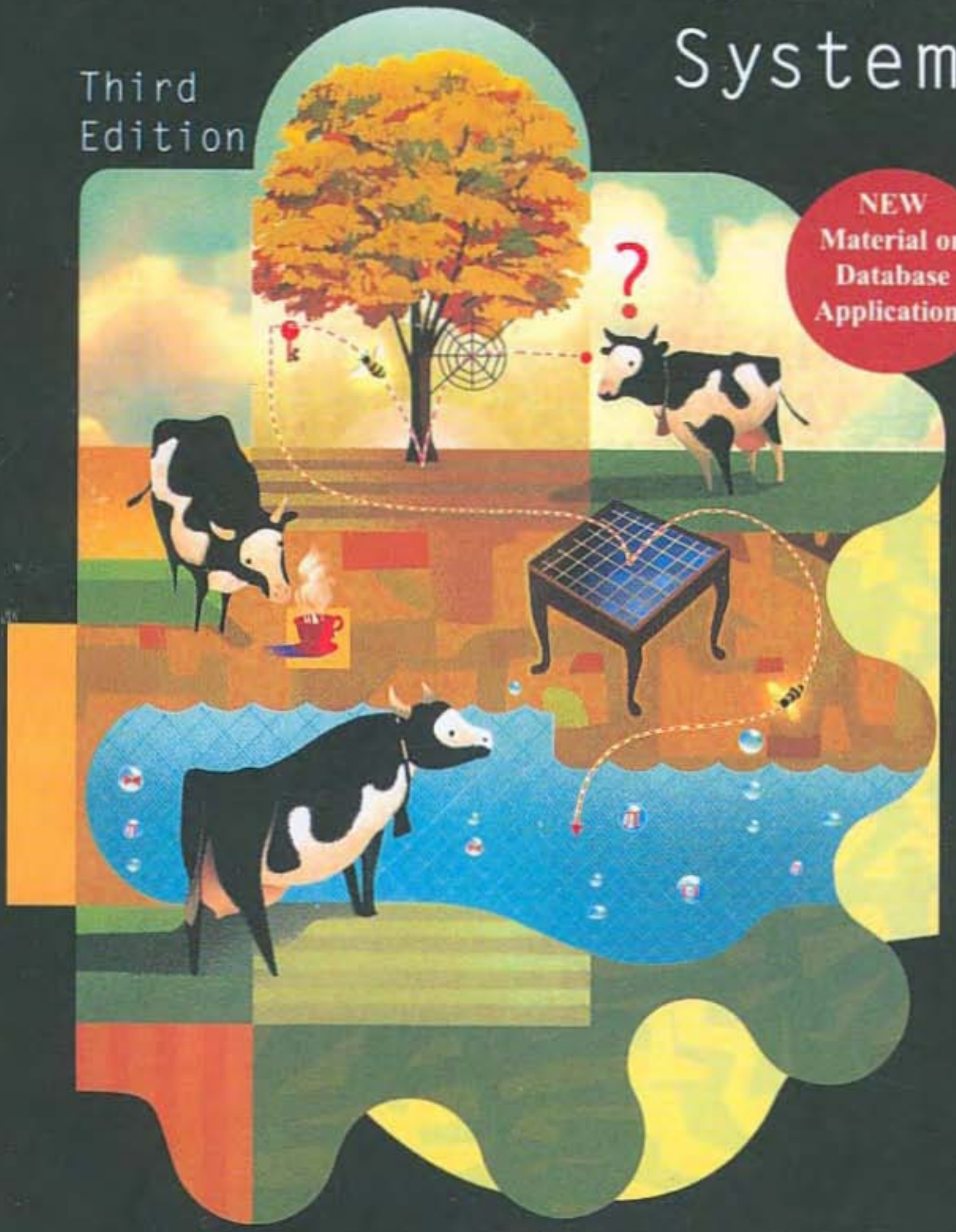
Database Management Systems

Third Edition



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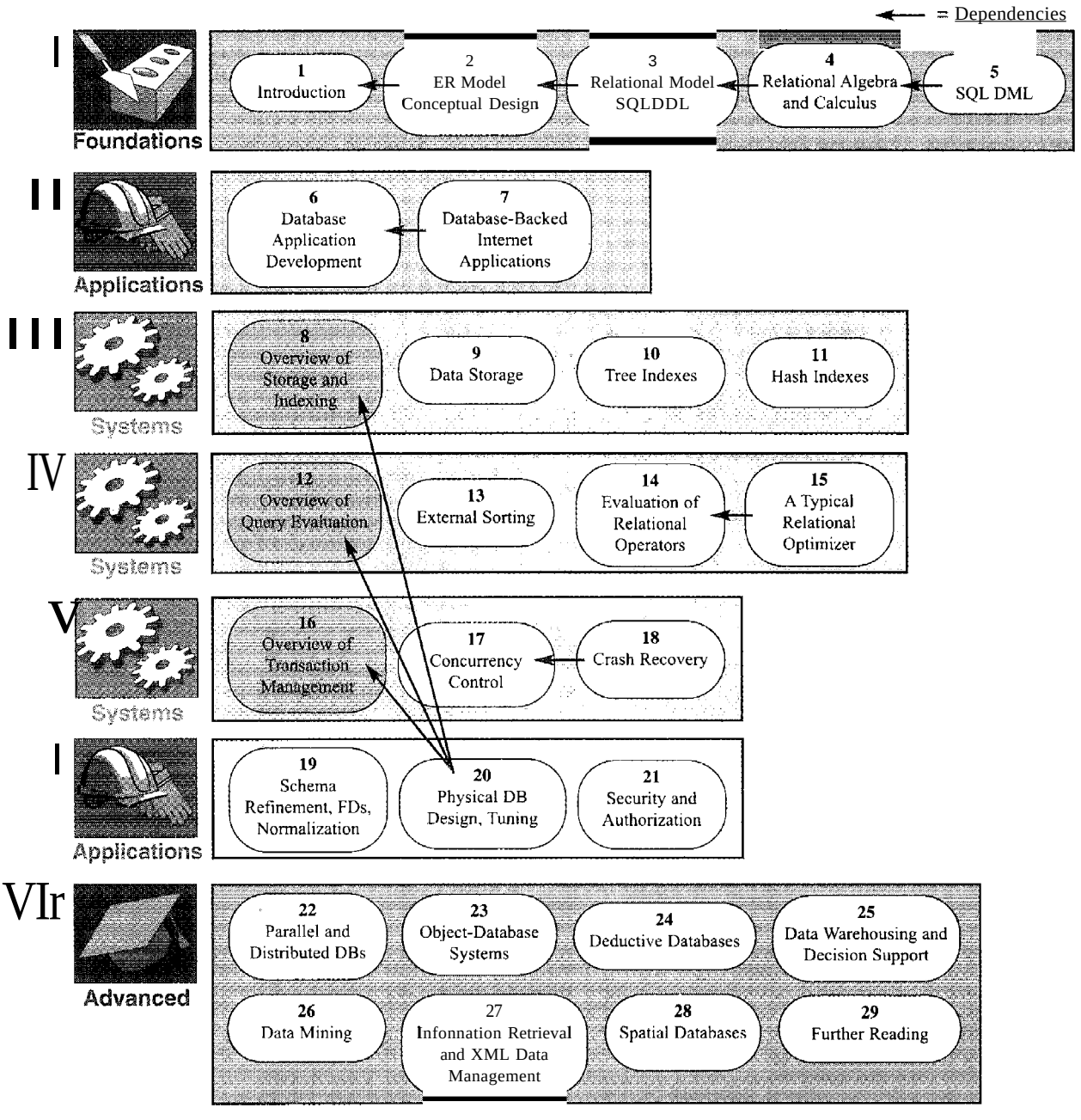

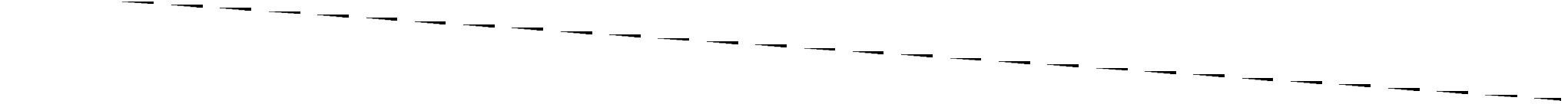

DATABASE MANAGEMENT SYSTEMS

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Third Edition

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To Apu, Ketan, and Vivek with love

To Keiko and Elisa

CONTENTS

PREFACE

> The advantage of doing one's praising for oneself is that one can lay it on so thick and exactly in the right places.
>
> --Samuel Butler

Database management systems are now an indispensable tool for managing information, and a course on the principles and practice of database systems is now an integral part of computer science curricula. This book covers the fundamentals of modern database management systems, in particular relational database systems.

We have attempted to present the material in a clear, simple style. A quantitative approach is used throughout with many detailed examples. An extensive set of exercises (for which solutions are available online to instructors) accompanies each chapter and reinforces students' ability to apply the concepts to real problems.

The book can be used with the accompanying software and programming assignments in two distinct kinds of introductory courses:

1. **Applications Emphasis:** A course that covers the principles of database systems, and emphasizes how they are used in developing data-intensive applications. Two new chapters on application development (one on database-backed applications, and one on Java and Internet application architectures) have been added to the third edition, and the entire book has been extensively revised and reorganized to support such a course. A running case-study and extensive online materials (e.g., code for SQL queries and Java applications, online databases and solutions) make it easy to teach a hands-on application-centric course.

2. Systems **Emphasis:** A course that has a strong systems emphasis and assumes that students have good programming skills in C and *C++*. In this case the accompanying Minibase software can be llsed as the basis for projects in which students are asked to implement various parts of a relational DBMS. Several central modules in the project software (e.g., heap files, buffer manager, B+ trees, hash indexes, various join methods)

are described in sufficient detail in the text to enable students to implement them, given the (C++) class interfaces.

Many instructors will no doubt teach a course that falls between these two extremes. The restructuring in the third edition offers a very modular organization that facilitates such hybrid courses. The also book contains enough material to support advanced courses in a two-course sequence.

Organization of the Third Edition

The book is organized into six main parts plus a collection of advanced topics, as shown in Figure 0.1. The Foundations chapters introduce database systems, the

(1) Foundations	Both
(2) Application Development	Applications emphasis
(3) Storage and Indexing	Systems emphasis
(4) Query Evaluation	Systems emphasis
(5) Transaction Management	Systems emphasis
(6) Database Design and Tuning	Applications emphasis
(7) Additional Topics	Both

Figure 0.1 Organization of Parts in the Third Edition

ER model and the relational model. They explain how databases are created and used, and cover the basics of database design and querying, including an in-depth treatment of SQL queries. While an instructor can omit some of this material at their discretion (e.g., relational calculus, some sections on the ER model or SQL queries), this material is relevant to every student of database systems, and we recommend that it be covered in as much detail as possible.

Each of the remaining five main parts has either an application or a systems empha.sis. Each of the three Systems parts has an overview chapter, designed to provide a self-contained treatment, e.g., Chapter 8 is an overview of storage and indexing. The overview chapters can be used to provide stand-alone coverage of the topic, or as the first chapter in a more detailed treatment. Thus, in an application-oriented course, Chapter 8 might be the only material covered on file organizations and indexing, whereas in a systems-oriented course it would be supplemented by a selection from Chapters 9 through 11. The Database Design and Tuning part contains a discussion of performance tuning and designing for secure access. These application topics are best covered after giving students a good grasp of database system architecture, and are therefore placed later in the chapter sequence.

Suggested Course Outlines

The book can be used in two kinds of introductory database courses, one with an applications emphasis and one with a systems emphasis.

The *introductory applications-oriented course* could cover the :Foundations chapters, then the Application Development chapters, followed by the overview systems chapters, and conclude with the Database Design and Tuning material. Chapter dependencies have been kept to a minimum, enabling instructors to easily fine tune what material to include. The Foundations material, Part I, should be covered first, and within Parts III, IV, and V, the overview chapters should be covered first. The only remaining dependencies between chapters in Parts I to **VI** are shown as arrows in Figure 0.2. The chapters in Part I should be covered in sequence. However, the coverage of algebra and calculus can be skipped in order to get to SQL queries sooner (although we believe this material is important and recommend that it should be covered before SQL).

The *introductory systems-oriented course* would cover the Foundations chapters and a selection of Applications and Systems chapters. An important point for systems-oriented courses is that the timing of programming projects (e.g., using Minibase) makes it desirable to cover some systems topics early. Chapter dependencies have been carefully limited to allow the Systems chapters to be covered as soon as Chapters 1 and 3 have been covered. The remaining Foundations chapters and Applications chapters can be covered subsequently.

The book also has ample material to support a multi-course sequence. Obviously, choosing an applications or systems emphasis in the introductory course results in dropping certain material from the course; the material in the book supports a comprehensive two-course sequence that covers both applications and systems aspects. The Additional Topics range over a broad set of issues, and can be used as the core material for an advanced course, supplemented with further readings.

Supplementary Material

This book comes with extensive online supplements:

- **Online Chapter:** To make space for new material such as application development, information retrieval, and XML, we've moved the coverage of QBE to an online chapter. Students can freely download the chapter from the book's web site, and solutions to exercises from this chapter are included in solutions manual.

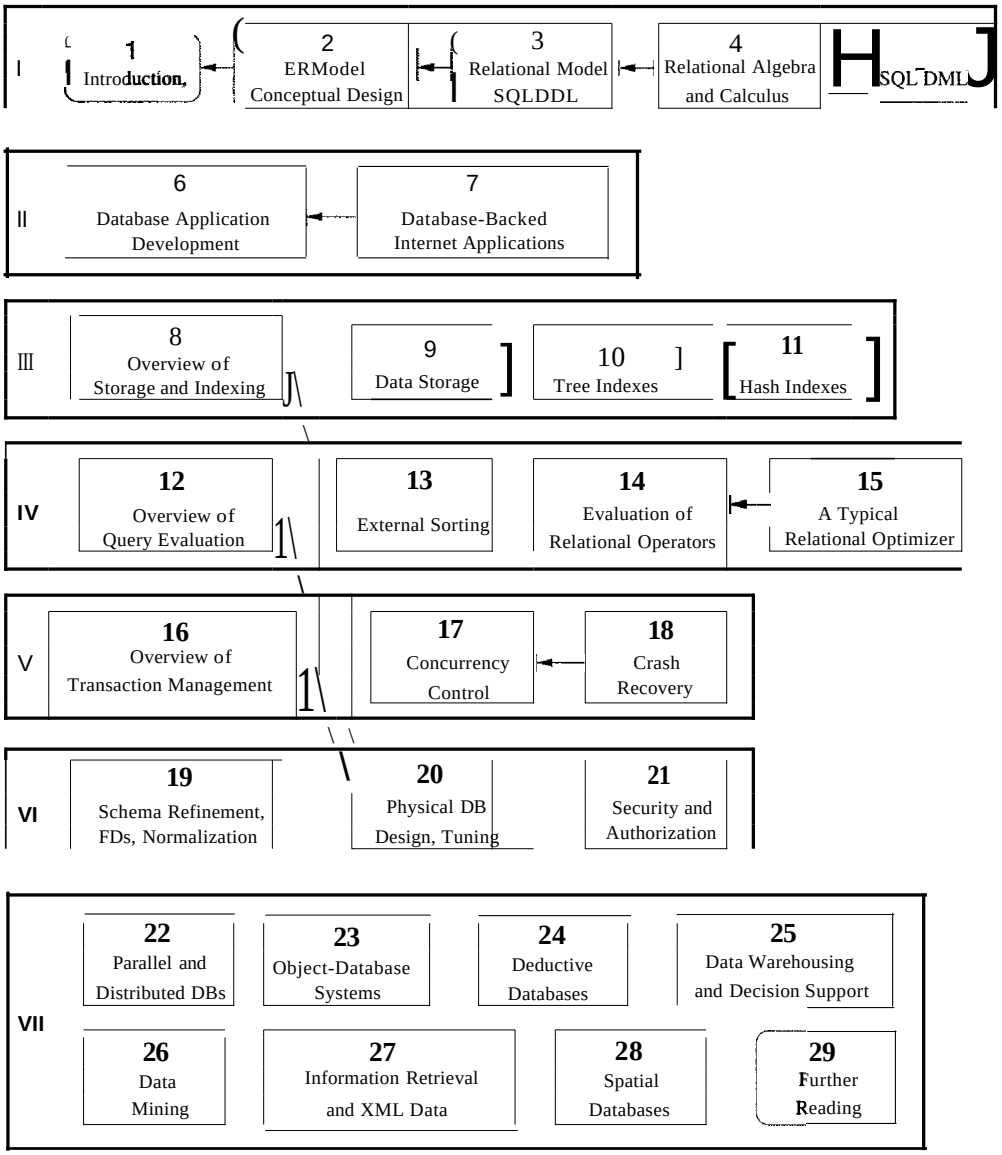

Figure 0.2 Chapter Organization and Dependencies

- Lecture Slides: Lecture slides are freely available for all chapters in Postscript, and PDF formats. Course instructors can also obtain these slides in Microsoft Powerpoint format, and can adapt them to their teaching needs. Instructors also have access to all figures llsed in the book (in xfig format), and can use them to modify the slides.

- Solutions to Chapter Exercises: The book has an unusually extensive set of in-depth exercises. Students can obtain solutioIls to odd-numbered chapter exercises and a set of lecture slides for each chapter through the Web in Postscript and Adobe PDF formats. Course instructors can obtain solutions to all exercises.
- Software: The book comes with two kinds of software. First, we have Minibase, a small relational DBMS intended for use in systems-oriented courses. Minibase comes with sample assignments and solutions, as described in Appendix 30. Access is restricted to course instructors. Second, we offer code for all SQL and Java application development exercises in the book, together with scripts to create sample databases, and scripts for setting up several commercial DBMSs. Students can only access solution code for odd-numbered exercises, whereas instructors have access to all solutions.
- Instructor's Manual: The book comes with an online manual that offers instructors comments on the material in each chapter. It provides a summary of each chapter and identifies choices for material to emphasize or omit. The manual also discusses the on-line supporting material for that chapter and offers numerous suggestions for hands-on exercises and projects. Finally, it includes samples of examination papers from courses taught by the authors using the book. It is restricted to course instructors.

For More Information

The home page for this book is at URL:

http://www.cs.wisc.edu/-dbbook

It contains a list of the changes between the 2nd and 3rd editions, and a frequently updated *link to all known errors in the book and its accompanying supplements.* Instructors should visit this site periodically or register at this site to be notified of important changes by email.

Acknowledgments

This book grew out of lecture notes for CS564, the introductory (senior/graduate level) database course at UvV-Madison. David De\Vitt developed this course and the Minirel project, in which students wrote several well-chosen parts of a relational DBMS. My thinking about this material was shaped by teaching CS564, and Minirel was the inspiration for Minibase, which is more comprehensive (e.g., it has a query optimizer and includes visualization software) but

tries to retain the spirit of MinireL Mike Carey and I jointly designed much of Minibase. My lecture notes (and in turn this book) were influenced by Mike's lecture notes and by Yannis Ioannidis's lecture slides.

Joe Hellerstein used the beta edition of the book at Berkeley and provided invaluable feedback, assistance on slides, and hilarious quotes. vVriting the chapter on object-database systems with Joe was a lot of fun.

C. Mohan provided invaluable assistance, patiently answering a number of questions about implementation techniques used in various commercial systems, in particular indexing, concurrency control, and recovery algorithms. Moshe Zloof answered numerous questions about QBE semantics and commercial systems based on QBE. Ron Fagin, Krishna Kulkarni, Len Shapiro, Jim Melton, Dennis Shasha, and Dirk Van Gucht reviewed the book and provided detailed feedback, greatly improving the content and presentation. Michael Goldweber at Beloit College, Matthew Haines at Wyoming, Michael Kifer at SUNY StonyBrook, Jeff Naughton at Wisconsin, Praveen Seshadri at Cornell, and Stan Zdonik at Brown also used the beta edition in their database courses and offered feedback and bug reports. In particular, Michael Kifer pointed out an error in the (old) algorithm for computing a minimal cover and suggested covering some SQL features in Chapter 2 to improve modularity. Gio Wiederhold's bibliography, converted to Latex format by S. Sudarshan, and Michael Ley's online bibliography on databases and logic programming were a great help while compiling the chapter bibliographies. Shaun Flisakowski and Uri Shaft helped me frequently in my never-ending battles with Latex.

lowe a special thanks to the many, many students who have contributed to the Minibase software. Emmanuel Ackaouy, Jim Pruyne, Lee Schumacher, and Michael Lee worked with me when I developed the first version of Minibase (much of which was subsequently discarded, but which influenced the next version). Emmanuel Ackaouy and Bryan So were my TAs when I taught CS564 using this version and went well beyond the limits of a TAship in their efforts to refine the project. Paul Aoki struggled with a version of Minibase and offered lots of useful eomments as a TA at Berkeley. An entire class of CS764 students (our graduate database course) developed much of the current version of Minibase in a large class project that was led and coordinated by Mike Carey and me. Amit Shukla and Michael Lee were my TAs when I first taught CS564 using this version of Minibase and developed the software further.

Several students worked with me on independent projects, over a long period of time, to develop Minibase components. These include visualization packages for the buffer manager and B+ trees (Huseyin Bektas, Harry Stavropoulos, and Weiqing Huang); a query optimizer and visualizer (Stephen Harris, Michael Lee, and Donko Donjerkovic); an ER diagram tool based on the Opossum schema

editor (Eben Haber); and a GUI-based tool for normalization (Andrew Prock and Andy Therber). In addition, Bill Kimmel worked to integrate and fix a large body of code (storage manager, buffer manager, files and access methods, relational operators, and the query plan executor) produced by the CS764 class project. Ranjani Ramamurty considerably extended Bill's work on cleaning up and integrating the various modules. Luke Blanshard, Uri Shaft, and Shaun Flisakowski worked on putting together the release version of the code and developed test suites and exercises based on the Minibase software. Krishna Kunchithapadam tested the optimizer and developed part of the Minibase GUI.

Clearly, the Minibase software would not exist without the contributions of a great many talented people. With this software available freely in the public domain, I hope that more instructors will be able to teach a systems-oriented database course with a blend of implementation and experimentation to complement the lecture material.

I'd like to thank the many students who helped in developing and checking the solutions to the exercises and provided useful feedback on draft versions of the book. In alphabetical order: X. Bao, S. Biao, M. Chakrabarti, C. Chan, W. Chen, N. Cheung, D. Colwell, C. Fritz, V. Ganti, J. Gehrke, G. Glass, V. Gopalakrishnan, M. Higgins, T. Jasmin, M. Krishnaprasad, Y. Lin, C. Liu, M. Lusignan, H. Modi, S. Narayanan, D. Randolph, A. Ranganathan, J. Reminga, A. Therber, M. Thomas, Q. Wang, R. Wang, Z. Wang, and J. Yuan. Arcady GrenadeI', James Harrington, and Martin Reames at Wisconsin and Nina Tang at Berkeley provided especially detailed feedback.

Charlie Fischer, Avi Silberschatz, and Jeff Ullman gave me invaluable advice on working with a publisher. My editors at McGraw-Hill, Betsy Jones and Eric Munson, obtained extensive reviews and guided this book in its early stages. Emily Gray and Brad Kosirog were there whenever problems cropped up. At Wisconsin, Ginny Werner really helped me to stay on top of things.

Finally, this book was a thief of time, and in many ways it was harder on my family than on me. My sons expressed themselves forthrightly. From my (then) five-year-old, Ketan: "Dad, stop working on that silly book. You don't have any time for *me."* Two-year-old Vivek: "You working *boook?* No no no come play basketball me!" All the seasons of their discontent were visited upon my wife, and Apu nonetheless cheerfully kept the family going in its usual chaotic, happy way all the many evenings and weekends I was wrapped up in this book. (Not to mention the days when I was wrapped up in being a faculty member!) As in all things, I can trace my parents' hand in much of this; my father, with his love of learning, and my mother, with her love of us, shaped me. My brother Kartik's contributions to this book consisted chiefly of phone calls in which he kept me from working, but if I don't acknowledge him, he's liable to

be annoyed. I'd like to thank my family for being there and giving meaning to everything I do. (There! I knew I'd find a legitimate reason to thank Kartik.)

Acknowledgments for the Second Edition

Emily Gray and Betsy Jones at McGraw-Hill obtained extensive reviews and provided guidance and support as we prepared the second edition. Jonathan Goldstein helped with the bibliography for spatial databases. The following reviewers provided valuable feedback on content and organization: Liming Cai at Ohio University, Costas Tsatsoulis at University of Kansas, Kwok-Bun Vue at University of Houston, Clear Lake, William Grosky at Wayne State University, Sang H. Son at University of Virginia, James M. Slack at Minnesota State University, Mankato, Herman Balsters at University of Twente, Netherlands, Karen C. Davis at University of Cincinnati, Joachim Hammer at University of Florida, Fred Petry at Tulane University, Gregory Speegle at Baylor University, Salih Yurttas at Texas A&M University, and David Chao at San Francisco State University.

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A special thanks to the many people who answered a detailed survey about how commercial systems support various features: At IBM, Mike Carey, Bruce Lindsay, C. Mohan, and James Teng; at Informix, M. Muralikrishna and Michael Ubell; at Microsoft, David Campbell, Goetz Graefe, and Peter Spiro; at Oracle, Hakan Jacobsson, Jonathan D. Klein, Muralidhar Krishnaprasad, and M. Ziauddin; and at Sybase, Marc Chanliau, Lucien Dimino, Sangeeta Doraiswamy, Hanuma Kodavalla, Roger MacNicol, and Tirumanjanam Rengarajan.

After reading about himself in the acknowledgment to the first edition, Ketan (now 8) had a simple question: "How come you didn't dedicate the book to us? Why mom?" Ketan, I took care of this inexplicable oversight. Vivek (now 5) was more concerned about the extent of his fame: "Daddy, is my name in *evvy* copy of your book? Do they have it in *evvy* compooter science department in the world'?" Vivek, I hope so. Finally, this revision would not have made it without Apu's and Keiko's support.

Acknowledgments for the Third Edition

We thank Raghav Kaushik for his contribution to the discussion of XML, and Alex Thomasian for his contribution to the coverage of concurrency control. A special thanks to Jim JVlelton for giving us a pre-publication copy of his book on object-oriented extensions in the SQL: 1999 standard, and catching several bugs in a draft of this edition. Marti Hearst at Berkeley generously permitted us to adapt some of her slides on Information Retrieval, and Alon Levy and Dan Sueiu were kind enough to let us adapt some of their lectures on XML. Mike Carey offered input on Web services.

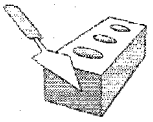
Emily Lupash at McGraw-Hill has been a source of constant support and encouragement. She coordinated extensive reviews from Ming Wang at Embry-Riddle Aeronautical University, Cheng Hsu at RPI, Paul Bergstein at Univ. of Massachusetts, Archana Sathaye at SJSU, Bharat Bhargava at Purdue, John Fendrich at Bradley, Ahmet Ugur at Central Michigan, Richard Osborne at Univ. of Colorado, Akira Kawaguchi at CCNY, Mark Last at Ben Gurion, Vassilis Tsotras at Univ. of California, and Ronald Eaglin at Univ. of Central Florida. It is a pleasure to acknowledge the thoughtful input we received from the reviewers, which greatly improved the design and content of this edition. Gloria Schiesl and Jade Moran dealt cheerfully and efficiently with last-minute snafus, and, with Sherry Kane, made a very tight schedule possible. Michelle Whitaker iterated many times on the cover and end-sheet design.

On a personal note for Raghu, Ketan, following the canny example of the camel that shared a tent, observed that "it is only fair" that Raghu dedicate this edition solely to him and Vivek, since "mommy already had it dedicated only to her." Despite this blatant attempt to hog the limelight, enthusiastically supported by Vivek and viewed with the indulgent affection of a doting father, this book is also dedicated to Apu, for being there through it all.

For Johannes, this revision would not have made it without Keiko's support and inspiration and the motivation from looking at Elisa's peacefully sleeping face.

PART I

FOUNDATIONS



1

OVERVIEW OF **DATABASE SYSTEMS**

- What is a DBMS, in particular, a relational DBMS?
- Why should we consider a DBMS to manage data?
- How is application data represented in a DBMS?
- How is data in a DBMS retrieved and manipulated?
- How does a DBMS support concurrent access and protect data during system failures?
- What are the main components of a DBMS?
- Who is involved with databases in real life?
- Key **concepts:** database management, data independence, database design, data model; relational databases and queries; schemas, levels of abstraction; transactions, concurrency and locking, recovery and logging; DBMS architecture; database administrator, application programmer, end user

Has everyone noticed that all the letters of the word *database* are typed with the left hand? Now the layout of the QWEHTY typewriter keyboard was designed, among other things, to facilitate the even use of both hands. It follows, therefore, that writing about databases is not only unnatural, but a lot harder than it appears.

---Anonymous

The alIlount of information available to us is literally exploding, and the value of data as an organizational asset is widely recognized. To get the most out of their large and complex datasets, users require tools that simplify the tasks of

The area of database management systenls is a microcosm of computer science in general. The issues addressed and the techniques used span a wide spectrum, including languages, object-orientation and other progTamming paradigms, compilation, operating systems, concurrent programming, data structures, algorithms, theory, parallel and distributed systems, user interfaces, expert systems and artificial intelligence, statistical techniques, and dynamic programming. We cannot go into all these aspects of database management in one book, but we hope to give the reader a sense of the excitement in this rich and vibrant discipline.

managing the data and extracting useful information in a timely fashion. Otherwise, data can become a liability, with the cost of acquiring it and managing it far exceeding the value derived from it.

A database is a collection of data, typically describing the activities of one or more related organizations. For example, a university database might contain information about the following:

- *Entities* such as students, faculty, courses, and classrooms.
- *Relationships* between entities, such as students' enrollment in courses, faculty teaching courses, and the use of rooms for courses.

A database management system, or DBMS, is software designed to assist in maintaining and utilizing large collections of data. The need for such systems, as well as their use, is growing rapidly. The alternative to using a DBMS is to store the data in files and write application-specific code to manage it. The use of a DBMS has several important advantages, as we will see in Section 1.4.

1.1 MANAGING DATA

The goal of this book is to present an in-depth introduction to database management systems, with an emphasis on how to *design* a database and *use* a DBMS effectively. Not surprisingly, many decisions about how to use a DBIvIS for a given application depend on what capabilities the DBMS supports efficiently. Therefore, to use a DBMS well, it is necessary to also understand how a DBMS *works*.

Many kinds of database management systems are in use, but this book concentrates on relational database systems (RDBMSs), which are by far the dominant type of DBMS today. The following questions are addressed in the corc chapters of this hook:

1. Database Design and Application Development: How can a user describe a real-world enterprise (e.g., a university) in terms of the data stored in a DBMS? \Vhat factors must be considered in deciding how to organize the stored data? How can ,ve develop applications that rely upon a DBMS? (Chapters 2, 3, 6, 7, 19, 20, and 21.)

2. Data Analysis: How can a user answer questions about the enterprise by posing queries over the data in the DBMS? (Chapters 4 and 5.)[1]

3. Concurrency and Robustness: How does a DBMS allow many users to access data concurrently, and how does it protect the data in the event of system failures? (Chapters 16, 17, and 18.)

4. Efficiency and Scalability: How does a DBMS store large datasets and answer questions against this data efficiently? (Chapters 8, 9, la, 11, 12, 13, 14, and 15.)

Later chapters cover important and rapidly evolving topics, such as parallel and distributed database management, data warehousing and complex queries for decision support, data mining, databases and information retrieval, XML repositories, object databases, spatial data management, and rule-oriented DBMS extensions.

In the rest of this chapter, we introduce these issues. In Section 1.2, we begin with a brief history of the field and a discussion of the role of database management in modern information systems. We then identify the benefits of storing data in a DBMS instead of a file system in Section 1.3, and discuss the advantages of using a DBMS to manage data in Section 1.4. In Section 1.5, we consider how information about an enterprise should be organized and stored in a DBMS. A user probably thinks about this information in high-level terms that correspond to the entities in the organization and their relationships, whereas the DBMS ultimately stores data in the form of (rnany, many) bits. The gap between how users think of their data and how the data is ultimately stored is bridged through several *levels of abstraction* supported by the DBMS. Intuitively, a user can begin by describing the data in fairly high-level terms, then refine this description by considering additional storage and representation details as needed.

In Section 1.6, we consider how users can retrieve data stored in a DBMS and the need for techniques to efficiently compute answers to questions involving such data. In Section 1.7, we provide an overview of how a DBMS supports concurrent access to data by several users and how it protects the data in the event of system failures.

[1] An online chapter on Query-by-Example (QBE) is also available.

We then briefly describe the internal structure of a DBMS in Section 1.8, and mention various groups of people associated with the development and use of a DBMS in Section 1.9.

1.2 A HISTORICAL PERSPECTIVE

From the earliest days of computers, storing and manipulating data have been a major application focus. The first general-purpose DBMS, designed by Charles Bachman at General Electric in the early 1960s, was called the Integrated Data Store. It formed the basis for the *network data model*, which was standardized by the Conference on Data Systems Languages (CODASYL) and strongly influenced database systems through the 1960s. Bachman was the first recipient of ACM's Turing Award (the computer science equivalent of a Nobel Prize) for work in the database area; he received the award in 1973.

In the late 1960s, IBM developed the Information Management System (IMS) DBMS, used even today in many major installations. IMS formed the basis for an alternative data representation framework called the *hierarchical data model*. The SABRE system for making airline reservations was jointly developed by American Airlines and IBM around the same time, and it allowed several people to access the same data through a computer network. Interestingly, today the same SABRE system is used to power popular Web-based travel services such as Travelocity.

In 1970, Edgar Codd, at IBM's San Jose Research Laboratory, proposed a new data representation framework called the *relational data model*. This proved to be a watershed in the development of database systems: It sparked the rapid development of several DBMSs based on the relational model, along with a rich body of theoretical results that placed the field on a firm foundation. Codd won the 1981 Turing Award for his seminal work. Database systems matured as an academic discipline, and the popularity of relational DBMSs changed the commercial landscape. Their benefits were widely recognized, and the use of DBMSs for managing corporate data became standard practice.

In the 1980s, the relational model consolidated its position as the dominant DBMS paradigm, and database systems continued to gain widespread use. The SQL query language for relational databases, developed as part of IBM's System R project, is now the standard query language. SQL was standardized in the late 1980s, and the current standard, SQL:1999, was adopted by the American National Standards Institute (ANSI) and International Organization for Standardization (ISO). Arguably, the most widely used form of concurrent programming is the concurrent execution of database programs (called *transactions*). Users write programs as if they are to be run by themselves, and

the responsibility for running them concurrently is given to the DBMS. James Gray won the 1999 Turing award for his contributions to database transaction management.

In the late 1980s and the 1990s, advances were made in many areas of database systems. Considerable research was carried out into more powerful query languages and richer data models, with emphasis placed on supporting complex analysis of data from all parts of an enterprise. Several vendors (e.g., IBM's DB2, Oracle 8, Informix[2] UDS) extended their systems with the ability to store new data types such as images and text, and to ask more complex queries. Specialized systems have been developed by numerous vendors for creating *data warehouses*, consolidating data from several databases, and for carrying out specialized analysis.

An interesting phenomenon is the emergence of several enterprise resource planning (ERP) and management resource planning (MRP) packages, which add a substantial layer of application-oriented features on top of a DBMS. Widely used packages include systems from Baan, Oracle, PeopleSoft, SAP, and Siebel. These packages identify a set of common tasks (e.g., inventory management, human resources planning, financial analysis) encountered by a large number of organizations and provide a general application layer to carry out these tasks. The data is stored in a relational DBMS and the application layer can be customized to different companies, leading to lower overall costs for the companies, compared to the cost of building the application layer from scratch.

Most significant, perhaps, DBMSs have entered the Internet Age. While the first generation of websites stored their data exclusively in operating systems files, the use of a DBMS to store data accessed through a Web browser is becoming widespread. Queries are generated through Web-accessible forms and answers are formatted using a markup language such as HTML to be easily displayed in a browser. All the database vendors are adding features to their DBMS aimed at making it more suitable for deployment over the Internet.

Database management continues to gain importance as more and more data is brought online and made ever more accessible through computer networking. Today the field is being driven by exciting visions such as multimedia databases, interactive video, streaming data, digital libraries, a host of scientific projects such as the human genome mapping effort and NASA's Earth Observation System project, and the desire of companies to consolidate their decision-making processes and *mine* their data repositories for useful information about their businesses. Commercially, database management systems represent one of the

2Informix was recently acquired by IBM.

largest and most vigorous market segments. Thus the study of database systems could prove to be richly rewarding in more ways than one!

1.3 FILE SYSTEMS VERSUS A DBMS

To understand the need for a DBMS, let us consider a motivating scenario: A company has a large collection (say, 500 GB[3]) of data on employees, departments, products, sales, and so on. This data is accessed concurrently by several employees. Questions about the data must be answered quickly, changes made to the data by different users must be applied consistently, and access to certain parts of the data (e.g., salaries) must be restricted.

We can try to manage the data by storing it in operating system files. This approach has many drawbacks, including the following:

- We probably do not have 500 GB of main memory to hold all the data. We must therefore store data in a storage device such as a disk or tape and bring relevant parts into main memory for processing as needed.
- Even if we have 500 GB of main memory, on computer systems with 32-bit addressing, we cannot refer directly to more than about 4 GB of data. We have to program some method of identifying all data items.
- We have to write special programs to answer each question a user may want to ask about the data. These programs are likely to be complex because of the large volume of data to be searched.
- We must protect the data from inconsistent changes made by different users accessing the data concurrently. If applications must address the details of such concurrent access, this adds greatly to their complexity.
- We must ensure that data is restored to a consistent state if the system crashes while changes are being made.
- Operating systems provide only a password mechanism for security. This is not sufficiently flexible to enforce security policies in which different users have permission to access different subsets of the data.

A DBMS is a piece of software designed to make the preceding tasks easier. By storing data in a DBMS rather than as a collection of operating system files, we can use the DBMS's features to manage the data in a robust and efficient rnanner. As the volume of data and the number of users grow hundreds of gigabytes of data and thousands of users are common in current corporate databases DBMS support becomes indispensable.

[3]A kilobyte (KB) is 1024 bytes, a megabyte (MB) is 1024 KBs, a gigabyte (GB) is 1024 MBs, a terabyte ('1'B) is 1024 CBs, and a petabyte (PB) is 1024 terabytes.

1.4 ADVANTAGES OF A DBMS

Using a DBMS to manage data has many advantages:

- Data Independence: Application programs should not, ideally, be exposed to details of data representation and storage, The DBMS provides an abstract view of the data that hides such details.
- Efficient Data Access: A DBMS utilizes a variety of sophisticated techniques to store and retrieve data efficiently. This feature is especially impOl'tant if the data is stored on external storage devices.
- Data Integrity and Security: If data is always accessed through the DBMS, the DBMS can enforce integrity constraints. For example, before inserting salary information for an employee, the DBMS can check that the department budget is not exceeded. Also, it can enforce *access controls* that govern what data is visible to different classes of users.
- Data Administration: When several users share the data, centralizing the administration of data can offer sig11ificant improvements. Experienced professionals who understand the nature of the data being managed, and how different groups of users use it, can be responsible for organizing the data representation to minimize redundancy and for fine-tuning the storage of the data to make retrieval efficient.
- Concurrent Access and Crash Recovery: A DBMS schedules concurrent accesses to the data in such a manner that users can think of the data as being accessed by only one user at a time. Further, the DBMS protects users from the effects of system failures.
- Reduced Application Development Time: Clearly, the DBMS supports important functions that are common to many applications accessing data in the DBMS. This, in conjunction with the high-level interface to the data, facilitates quick application development. DBMS applications are also likely to be more robust than similar stand-alone applications because many important tasks are handled by the DBMS (and do not have to be debugged and tested in the application).

Given all these advantages, is there ever a reason *not* to use a DBMS? Sometimes, yes. A DBMS is a complex piece of software, optimized for certain kinds of workloads (e.g., answering complex queries or handling many concurrent requests), and its performance may not be adequate for certain specialized applications. Examples include applications with tight real-time constraints or just a few well-defined critical operations for which efficient custom code must be written. Another reason for not using a DBMS is that an application may need to manipulate the data in ways not supported by the query language. In

such a situation, the abstract view of the data presented by the DBMS does not match the application's needs and actually gets in the way. As an example, relational databases do not support flexible analysis of text data (although vendors are now extending their products in this direction).

If specialized performance or data manipulation requirements are central to an application, the application may choose not to use a DBMS, especially if the added benefits of a DBMS (e.g., flexible querying, security, concurrent access, and crash recovery) are not required. In most situations calling for large-scale data management, however, DBMSs have become an indispensable tool.

1.5 DESCRIBING AND STORING DATA IN A DBMS

The user of a DBMS is ultimately concerned with some real-world enterprise, and the data to be stored describes various aspects of this enterprise. For example, there are students, faculty, and courses in a university, and the data in a university database describes these entities and their relationships.

A **data model** is a collection of high-level data description constructs that hide many low-level storage details. A DBMS allows a user to define the data to be stored in terms of a data model. Most database management systems today are based on the **relational data model,** which we focus on in this book.

While the data model of the DBMS hides many details, it is nonetheless closer to how the DBMS stores data than to how a user thinks about the underlying enterprise. A **semantic data model** is a more abstract, high-level data model that makes it easier for a user to come up with a good initial description of the data in an enterprise. These models contain a wide variety of constructs that help describe a real application scenario. A DBMS is not intended to support all these constructs directly; it is typically built around a data model with just a few basic constructs, such as the relational model. A database design in terms of a semantic model serves as a useful starting point and is subsequently translated into a database design in terms of the data model the DBMS actually supports.

A widely used semantic data model called the entity-relationship (ER) model allows us to pictorially denote entities and the relationships among them. We cover the ER model in Chapter 2.

An Example of Poor Design: The relational schema for Students illustrates a poor design choice; you should neVCT create a field such as *age*, whose value is constantly changing. A better choice would be *DOB* (for *date of birth*); age can be computed from this. \Ve continue to use *age* in our examples, however, because it makes them easier to read.

1.5.1 The Relational Model

In this section we provide a brief introduction to the relational model. The central data description construct in this model is a relation, which can be thought of as a set of records.

A description of data in terms of a data model is called a schema. In the relational model, the schema for a relation specifies its name, the name of each field (or attribute or column), and the type of each field. As an example, student information in a university database may be stored in a relation with the following schema:

Students(*sid:* string, *name:* string, *login:* string,
age: integer, *gpa:* real)

The preceding schema says that each record in the Students relation has five fields, with field names and types as indicated. An example instance of the Students relation appears in Figure 1.1.

sid	*name*	*login*	*age*	*gpa*
53666	Jones	jones@cs	18	3.4
53688	Smith	smith@ee	18	3.2
53650	Smith	smith@math	19	3.8
53831	Madayan	madayan@music	11	1.8
53832	Guldu	guldu@music	12	2.0

Figure 1.1 An Instance of the Students Relation

Each row in the Students relation is a record that describes a student. The description is not completeo----for example, the student's height is not included--- but is presumably adequate for the intended applications in the university database. Every row follows the schema of the Students relation. The schema call therefore be regarded as a template for describing a student.

We can make the description of a collection of students more precise by specifying integrity constraints, which are conditions that the records in a relation

must satisfy. For example, we could specify that every student has a unique *sid* value. Observe that we cannot capture this information by simply adding another field to the Students schema. Thus, the ability to specify uniqueness of the values in a field increases the accuracy with which we can describe our data. The expressiveness of the constructs available for specifying integrity constraints is an important aspect of a data model.

Other Data Models

In addition to the relational data model (which is used in numerous systems, including IBM's DB2, Informix, Oracle, Sybase, Microsoft's Access, FoxBase, Paradox, Tandem, and Teradata), other important data models include the hierarchical model (e.g., used in IBM's IMS DBMS), the network model (e.g., used in IDS and IDMS), the object-oriented model (e.g., used in Objectstore and Versant), and the object-relational model (e.g., used in DBMS products from IBM, Informix, ObjectStore, Oracle, Versant, and others). While many databases use the hierarchical and network models and systems based on the object-oriented and object-relational models are gaining acceptance in the marketplace, the dominant model today is the relational model.

In this book, we focus on the relational model because of its wide use and importance. Indeed, the object-relational model, which is gaining in popularity, is an effort to combine the best features of the relational and object-oriented models, and a good grasp of the relational model is necessary to understand object-relational concepts. (We discuss the object-oriented and object-relational models in Chapter 23.)

1.5.2 Levels of Abstraction in a DBMS

The data in a DBMS is described at three levels of abstraction, as illustrated in Figure 1.2. The database description consists of a schema at each of these three levels of abstraction: the *conceptual, physical,* and *external.*

A data definition language (DDL) is used to define the external and coneeptual schemas. We discuss the DDL facilities of the Inost widely used database language, SQL, in Chapter 3. All DBMS vendors also support SQL commands to describe aspects of the physical schema, but these commands are not part of the SQL language standard. Information about the conceptual, external, and physical schemas is stored in the system catalogs (Section 12.1). We discuss the three levels of abstraction in the rest of this section.

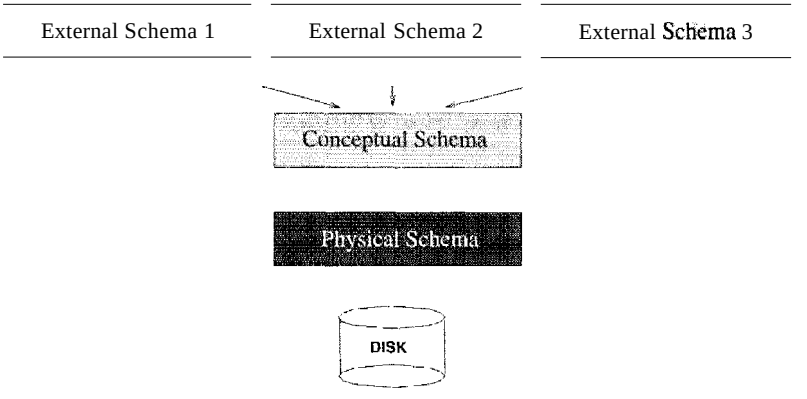


Figure 1.2 Levels of Abstraction in a DBMS

Conceptual Schema

The conceptual schema (sometimes called the logical schema) describes the stored data in terms of the data model of the DBMS. In a relational DBMS, the conceptual schema describes all relations that are stored in the database. In our sample university database, these relations contain information about *entities,* such as students and faculty, and about *relationships,* such as students' enrollment in courses. All student entities can be described using records in a Students relation, as we saw earlier. In fact, each collection of entities and each collection of relationships can be described as a relation, leading to the following conceptual schema:

Students(*sid:* string, *name:* string, *login:* string,
age: integer, *gpa:* real)
Faculty(*fid:* string, *fname:* string, *sal:* real)
Courses(*cid:* string, *cname:* string, *credits:* integer)
Rooms(*rno:* integer, *address:* string, *capacity:* integer)
Enrolled(*sid:* string, *cid:* string, *grade:* string)
Teaches(*fid:* string, *cid:* string)
Meets_In(*cid:* string, rno: integer, *time:* string)

The choice of relations, and the choice of fields for each relation, is not always obvious, and the process of arriving at a good conceptual schema is called conceptual database design. We discuss conceptual database design in Chapters 2 and 19.

Physical Schema

The physical schema specifies additional storage details. Essentially, the physical schema summarizes how the relations described in the conceptual schema are actually stored on secondary storage devices such as disks and tapes.

We must decide what file organizations to use to store the relations and create auxiliary data structures, called indexes, to speed up data retrieval operations. A sample physical schema for the university database follows:

- Store all relations as unsorted files of records. (A file in a DBMS is either a collection of records or a collection of pages, rather than a string of characters as in an operating system.)
- Create indexes on the first column of the Students, Faculty, and Courses relations, the *sal* column of Faculty, and the *capacity* column of Rooms.

Decisions about the physical schema are based on an understanding of how the data is typically accessed. The process of arriving at a good physical schema is called physical database design. We discuss physical database design in Chapter 20.

External Schema

External schemas, which usually are also in terms of the data model of the DBMS, allow data access to be customized (and authorized) at the level of individual users or groups of users. Any given database has exactly one conceptual schema and one physical schema because it has just one set of stored relations, but it may have several external schemas, each tailored to a particular group of users. Each external schema consists of a collection of one or more views and relations from the conceptual schema. A view is conceptually a relation, but the records in a view are not stored in the DBMS. Rather, they are computed using a definition for the view, in terms of relations stored in the DBMS. We discuss views in more detail in Chapters 3 and 25.

The external schema design is guided by end user requirements. For exalnple, we might want to allow students to find out the names of faculty members teaching courses as well as course enrollments. This can be done by defining the following view:

Courseinfo(*rid:* string, *fname:* string, *enrollment:* integer)

A user can treat a view just like a relation and ask questions about the records in the view. Even though the records in the view are not stored explicitly,

they are computed as needed. We did not include Courseinfo in the conceptual schema because we can compute Courseinfo from the relations in the conceptual schema, and to store it in addition would be redundant. Such redundancy, in addition to the wasted space, could lead to inconsistencies. For example, a tuple may be inserted into the Enrolled relation, indicating that a particular student has enrolled in some course, without incrementing the value in the *enrollment* field of the corresponding record of Courseinfo (if the latter also is part of the conceptual schema and its tuples are stored in the DBMS).

1.5.3 Data Independence

A very important advantage of using a DBMS is that it offers data independence. That is, application programs are insulated from changes in the way the data is structured and stored. Data independence is achieved through use of the three levels of data abstraction; in particular, the conceptual schema and the external schema provide distinct benefits in this area.

Relations in the external schema (view relations) are in principle generated on demand from the relations corresponding to the conceptual schema.[4] If the underlying data is reorganized, that is, the conceptual schema is changed, the definition of a view relation can be modified so that the same relation is computed as before. For example, suppose that the Faculty relation in our university database is replaced by the following two relations:

Faculty_public(*fid:* string, *fname:* string, *office:* integer)
Faculty_private(*fid:* string, *sal:* real)

Intuitively, some confidential information about faculty has been placed in a separate relation and information about offices has been added. The Courseinfo view relation can be redefined in terms of Faculty_public and Faculty_private, which together contain all the information in Faculty, so that a user who queries Courseinfo will get the same answers as before.

Thus, users can be shielded from changes in the logical structure of the data, or changes in the choice of relations to be stored. This property is called logical data independence.

In turn, the conceptual schema insulates users from changes in physical storage details. This property is referred to as physical data independence. The conceptual schema hides details such as how the data is actually laid out on disk, the file structure, and the choice of indexes. As long as the conceptual

4In practice, they could be precomputed and stored to speed up queries on view relations, but the computed view relations must be updated whenever the underlying relations are updated.

schema remains the same, we can change these storage details without altering applications. (Of course, performance might be affected by such changes.)

1.6 QUERIES IN A DBMS

The ease \vith which information can be obtained from a database often determines its value to a user. In contrast to older database systems, relational database systems allow a rich class of questions to be posed easily; this feature has contributed greatly to their popularity. Consider the sample university database in Section 1.5.2. Here are some questions a user might ask:

1. What is the name of the student with student ID 1234567
2. What is the average salary of professors who teach course CS5647
3. How many students are enrolled in CS5647
4. What fraction of students in CS564 received a grade better than B7
5. Is any student with a CPA less than 3.0 enrolled in CS5647

Such questions involving the data stored in a DBMS are called queries. A DBMS provides a specialized language, called the query language, in which queries can be posed. A very attractive feature of the relational model is that it supports powerful query languages. Relational calculus is a formal query language based on mathematical logic, and queries in this language have an intuitive, precise meaning. Relational algebra is another formal query language, based on a collection of operators for manipulating relations, which is equivalent in power to the calculus.

A DBMS takes great care to evaluate queries as efficiently as possible. We discuss query optimization and evaluation in Chapters 12, 14, and 15. Of course, the efficiency of query evaluation is determined to a large extent by how the data is stored physically. Indexes can be used to speed up many queries----in fact, a good choice of indexes for the underlying relations can speed up each query in the preceding list. We discuss data storage and indexing in Chapters 8, 9, 10, and 11.

A DBMS enables users to create, modify, and query data through a data manipulation language (DML). Thus, the query language is only one part of the Dl\ilL, which also provides constructs to insert, delete, and modify data,. We will discuss the DML features of SQL in Chapter 5. The DML and DDL are collectively referred to as the data sublanguage when embedded within a host language (e.g., C or COBOL).

1.7 TRANSACTION MANAGEMENT

Consider a database that holds information about airline reservations. At any given instant, it is possible (and likely) that several travel agents are looking up information about available seats on various flights and making new seat reservations. When several users access (and possibly modify) a database concurrently, the DBMS must order their requests carefully to avoid conflicts. For example, when one travel agent looks up Flight 100 on some given day and finds an empty seat, another travel agent may simultaneously be making a reservation for that seat, thereby making the information seen by the first agent obsolete.

Another example of concurrent use is a bank's database. While one user's application program is computing the total deposits, another application may transfer money from an account that the first application has just 'seen' to an account that has not yet been seen, thereby causing the total to appear larger than it should be. Clearly, such anomalies should not be allowed to occur. However, disallowing concurrent access can degrade performance.

Further, the DBMS must protect users from the effects of system failures by ensuring that all data (and the status of active applications) is restored to a consistent state when the system is restarted after a crash. For example, if a travel agent asks for a reservation to be made, and the DBMS responds saying that the reservation has been made, the reservation should not be lost if the system crashes. On the other hand, if the DBMS has not yet responded to the request, but is making the necessary changes to the data when the crash occurs, the partial changes should be undone when the system comes back up.

A transaction is *anyone execution* of a user program in a DBMS. (Executing the same program several times will generate several transactions.) This is the basic unit of change as seen by the DBMS: Partial transactions are not allowed, and the effect of a group of transactions is equivalent to some serial execution of all transactions. We briefly outline how these properties are guaranteed, deferring a detailed discussion to later chapters.

1.7.1 Concurrent Execution of **Transactions**

An important task of a DBMS is to schedule concurrent accesses to data so that each user can safely ignore the fact that others are accessing the data concurrently. The importance of this task cannot be underestimated because a database is typically shared by a large number of users, who submit their requests to the DBMS independently and simply cannot be expected to deal with arbitrary changes being made concurrently by other users. A DBMS

allows users to think of their programs as if they were executing in isolation, one after the other in some order chosen by the DBMS. For example, if a progTam that deposits cash into an account is submitted to the DBMS at the same time as another program that debits money from the same account, either of these programs could be run first by the DBMS, but their steps will not be interleaved in such a way that they interfere with each other.

A locking protocol is a set of rules to be followed by each transaction (and enforced by the DBMS) to ensure that, even though actions of several transactions might be interleaved, the net effect is identical to executing all transactions in some serial order. A lock is a mechanism used to control access to database objects. Two kinds of locks are commonly supported by a DBMS: shared locks on an object can be held by two different transactions at the same time, but an exclusive lock on an object ensures that no other transactions hold *any* lock on this object.

Suppose that the following locking protocol is followed: *Every transaction begins by obtaining a shared lock on each data object that it needs to read and an exclusive lock on each data object that it needs to modify, then releases all its locks after completing all actions.* Consider two transactions *T1* and *T2* such that *T1* wants to modify a data object and *T2* wants to read the same object. Intuitively, if *T1's* request for an exclusive lock on the object is granted first, *T2* cannot proceed until *T1* releases this lock, because *T2's* request for a shared lock will not be granted by the DBMS until then. Thus, all of *T1's* actions will be completed before any of *T2's* actions are initiated. We consider locking in more detail in Chapters 16 and 17.

1.7.2 Incomplete Transactions and System Crashes

Transactions can be interrupted before running to completion for a va,riety of reasons, e.g., a system crash. A DBMS must ensure that the changes made by such incomplete transactions are removed from the database. For example, if the DBMS is in the middle of transferring money from account A to account B and has debited the first account but not yet credited the second when the crash occurs, the money debited from account A must be restored when the system comes back up after the crash.

To do so, the DBMS maintains a log of all writes to the database. A crucial property of the log is that each write action must be recorded in the log (on disk) *before* the corresponding change is reflected in the database itself--otherwise, if the system crashes just after making the change in the database but before the change is recorded in the log, the DBIVIS would be unable to detect and undo this change. This property is called Write-Ahead Log, or WAL. To ensure

this property, the DBMS must be able to selectively force a page in memory to disk.

The log is also used to ensure that the changes made by a successfully completed transaction are not lost due to a system crash, as explained in Chapter 18. Bringing the database to a consistent state after a system crash can be a slow process, since the DBMS must ensure that the effects of all transactions that completed prior to the crash are restored, and that the effects of incomplete transactions are undone. The time required to recover from a crash can be reduced by periodically forcing some information to disk; this periodic operation is called a checkpoint.

1.7.3 Points to Note

In summary, there are three points to remember with respect to DBMS support for concurrency control and recovery:

1. Every object that is read or written by a transaction is first locked in shared or exclusive mode, respectively. Placing a lock on an object restricts its availability to other transactions and thereby affects performance.
2. For efficient log maintenance, the DBMS must be able to selectively force a collection of pages in main memory to disk. Operating system support for this operation is not always satisfactory.
3. Periodic checkpointing can reduce the time needed to recover from a crash. Of course, this must be balanced against the fact that checkpointing too often slows down normal execution.

1.8 STRUCTURE OF A DBMS

Figure 1.3 shows the structure (with some simplification) of a typical DBMS based on the relational data model.

The DBMS accepts SQL comma,nels generated from a variety of user interfaces, produces query evaluation plans, executes these plans against the database, and returns the answers. (This is a simplification: SQL commands can be embedded in host-language application programs, e.g., Java or COBOL programs. We ignore these issues to concentrate on the core DBMS functionality.)

When a user issues a query, the parsed query is presented to a query optimizer, which uses information about how the data is stored to produce an efficient execution plan for evaluating the query. An execution plan is a

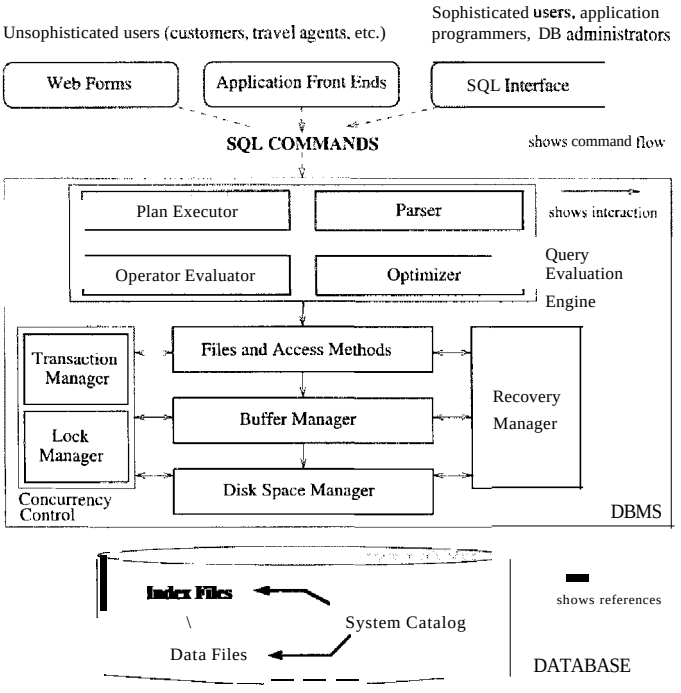


Figure 1.3 Architecture of a DBMS

blueprint for evaluating a query, usually represented as a tree of relational operators (with annotations that contain additional detailed information about which access methods to use, etc.). We discuss query optimization in Chapters 12 and 15. Relational operators serve as the building blocks for evaluating queries posed against the data. The implementation of these operators is discussed in Chapters 12 and 14.

The code that implements relational operators sits on top of the file and access methods layer. This layer supports the concept of a file, which, in a DBMS, is a collection of pages or a collection of records. Heap files, or files of unordered pages, as well as indexes are supported. In addition to keeping track of the pages in a file, this layer organizes the information within a page. File and page level storage issues are considered in Chapter 9. File organizations and indexes are cQIlsidered in Chapter 8.

The files and access methods layer code sits on top of the buffer manager, which brings pages in from disk to main memory as needed in response to read requests. Buffer management is discussed in Chapter 9.

The lowest layer of the DBMS software deals with management of space on disk, where the data is stored. Higher layers allocate, deallocate, read, and write pages through (routines provided by) this layer, called the disk space manager. This layer is discussed in Chapter 9.

The DBMS supports concurrency and crash recovery by carefully scheduling user requests and maintaining a log of all changes to the database. DBMS components associated with concurrency control and recovery include the transaction manager, which ensures that transactions request and release locks according to a suitable locking protocol and schedules the execution transactions; the lock manager, which keeps track of requests for locks and grants locks on database objects when they become available; and the recovery manager, which is responsible for maintaining a log and restoring the system to a consistent state after a crash. The disk space manager, buffer manager, and file and access method layers must interact with these components. We discuss concurrency control and recovery in detail in Chapter 16.

1.9 PEOPLE WHO WORK WITH DATABASES

Quite a variety of people are associated with the creation and use of databases. Obviously, there are database implementors, who build DBMS software, and end users who wish to store and use data in a DBMS. Database implementors work for vendors such as IBM or Oracle. End users come from a diverse and increasing number of fields. As data grows in complexity and volume, and is increasingly recognized as a major asset, the importance of maintaining it professionally in a DBMS is being widely accepted. Many end users simply use applications written by database application programmers (see below) and so require little technical knowledge about DBMS software. Of course, sophisticated users who make more extensive use of a DBMS, such as writing their own queries, require a deeper understanding of its features.

In addition to end users and implementors, two other classes of people are associated with a DBMS: *application programmers* and *database administrators*.

Database application programmers develop packages that facilitate data access for end users, who are usually not computer professionals, using the host or data languages and software tools that DBMS vendors provide. (Such tools include report writers, spreadsheets, statistical packages, and the like.) Application programs should ideally access data through the external schema. It is possible to write applications that access data at a lower level, but such applications would comprornise data independence.

A personal database is typically maintained by the individual who owns it and uses it. However, corporate or enterprise-wide databases are typically important enough and complex enough that the task of designing and maintaining the database is entrusted to a professional, called the database administrator (DBA). The DBA is responsible for many critical tasks:

- Design of the Conceptual and Physical Schemas: The DBA is responsible for interacting with the users of the system to understand what data is to be stored in the DBMS and how it is likely to be used. Based on this knowledge, the DBA must design the conceptual schema (decide what relations to store) and the physical schema (decide how to store them). The DBA may also design widely used portions of the external schema, although users probably augment this schema by creating additional views.
- Security and Authorization: The DBA is responsible for ensuring that unauthorized data access is not permitted. In general, not everyone should be able to access all the data. In a relational DBMS, users can be granted permission to access only certain views and relations. For example, although you might allow students to find out course enrollments and who teaches a given course, you would not want students to see faculty salaries or each other's grade information. The DBA can enforce this policy by giving students permission to read only the Courseinfo view.
- Data Availability and Recovery from Failures: The DBA must take steps to ensure that if the system fails, users can continue to access as much of the uncorrupted data as possible. The DBA must also work to restore the data to a consistent state. The *DB.I\!IS* provides software support for these functions, but the DBA is responsible for implementing procedures to back up the data periodically and maintain logs of system activity (to facilitate recovery from a crash).
- Database Tuning: Users' needs are likely to evolve with time. The DBA is responsible for modifying the database, in particular the conceptual and physical schemas, to ensure adequate performance as requirements change.

1.10 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are the main benefits of using a DBMS to manage data in applications involving extensive data access? (Sections 1.1, 1.4)
- When would you store data in a DBMS instead of in operating system files and vice-versa? (Section 1.3)

- What is a data model? \Vhat is the relational data model? What is data independence and how does a DBMS support it? (Section 1.5)
- Explain the advantages of using a query language instead of custom programs to process data. (Section 1.6)
- What is a transaction? \Vhat guarantees does a DBMS offer with respect to transactions? (Section 1.7)
- What are locks in a DBMS, and why are they used? What is write-ahead logging, and why is it used? What is checkpointing and why is it used? (Section 1.7)
- Identify the main components in a DBMS and briefly explain what they do. (Section 1.8)
- Explain the different roles of database administrators, application programmers, and end users of a database. Who needs to know the most about database systems? (Section 1.9)

EXERCISES

Exercise 1.1 Why would you choose a database system instead of simply storing data in operating system files? When would it make sense *not* to use a database system?

Exercise 1.2 What is logical data independence and why is it important?

Exercise 1.3 Explain the difference between logical and physical data independence.

Exercise 1.4 Explain the difference between external, internal, and conceptual schemas. How are these different schema layers related to the concepts of logical and physical data independence?

Exercise 1.5 What are the responsibilities of a DBA? If we assume that the DBA is never interested in running his or her own queries, does the DBA still need to understand query optimization? Why?

Exercise 1.6 Scrooge McNugget wants to store information (names, addresses, descriptions of embarrassing moments, etc.) about the many ducks on his payroll. Not surprisingly, the volume of data compels him to buy a database system. To save money, he wants to buy one with the fewest possible features, and he plans to run it as a stand-alone application on his PC clone. Of course, Scrooge does not plan to share his list with anyone. Indicate which of the following DBMS features Scrooge should pay for; in each case, also indicate why Scrooge should (or should not) pay for that feature in the system he buys.

1. A security facility.
2. Concurrency control.
3. Crash recovery.
4. A view mechanism.

5. A query language.

Exercise 1.7 Which of the following plays an important role in *representing* information about the real world in a database'? Explain briefly.

1. The data definition language.
2. The data manipulation language.
3. The buffer manager.
4. The data model.

Exercise 1.8 Describe the structure of a DBMS. If your operating system is upgraded to support some new functions on as files (e.g., the ability to force some sequence of bytes to disk), which layer(s) of the DBMS would you have to rewrite to take advantage of these new functions?

Exercise 1.9 Answer the following questions:

1. What is a transaction?
2. Why does a DBMS interleave the actions of different transactions instead of executing transactions one after the other?
3. What must a user guarantee with respect to a transaction and database consistency? What should a DBMS guarantee with respect to concurrent execution of several transactions and database consistency'?
4. Explain the strict two-phase locking protocol.
5. What is the WAL property, and why is it important?

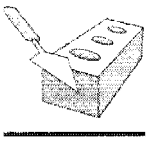
PROJECT-BASED EXERCISES

Exercise 1.10 Use a Web browser to look at the HTML documentation for Minibase. Try to get a feel for the overall architecture.

BIBLIOGRAPHIC NOTES

The evolution of database management systems is traced in [289]. The use of data models for describing real-world data is discussed in [423], and [425] contains a taxonomy of data models. The three levels of abstraction were introduced in [186, 712]. The network data model is described in [186], and [775] discusses several commercial systems based on this model. [721] contains a good annotated collection of systems-oriented papers on database management.

Other texts covering database management systems include [204, 245, 305, 339, 475, 574, 689, 747, 762]. [204] provides a detailed discussion of the relational model from a conceptual standpoint and is notable for its extensive annotated bibliography. [574] presents a performance-oriented perspective, with references to several commercial systems. [245] and [689] offer broad coverage of databa,se system concepts, including a discussion of the hierarchical and network data models. [339] emphasizes the connection between database query languages and logic programming. [762] emphasizes data models. Of these texts, [747] provides the most detailed discussion of theoretical issues. Texts devoted to theoretical aspects include [3, 45, 501]. Handbook [744] includes a section on databases that contains introductory survey articles on a number of topics.



2

INTRODUCTION TO DATABASE DESIGN

- What are the steps in designing a database?
- Why is the ER model used to create an initial design?
- What are the main concepts in the ER model?
- What are guidelines for using the ER model effectively?
- How does database design fit within the overall design framework for complex software within large enterprises?
- What is UML and how is it related to the ER model?
- Key concepts: database design, conceptual, logical, and physical design; entity-relationship (ER) model, entity set, relationship set, attribute, instance, key; integrity constraints, one-to-many and many-to-many relationships, participation constraints; weak entities, class hierarchies, aggregation; UML, class diagrams, database diagrams, component diagrams.

The great successful men of the world have used their imaginations. They think ahead and create their mental picture. and then go to work materializing that picture in all its details, filling in here, adding a little there, altering this bit and that bit, but steadily building, steadily building.

Robert Collier

The *entity-relationship (ER) data 'model* allows us to describe the data involved in a real-world enterprise in terms of objects and their relationships and is widely used to (levelop an initial database design. It provides useful eoncepts that allow us to move fronl an informal description of what users want 1'rom

their database to a more detailed, precise description that can be implemented in a DBMS. In this chapter, we introduce the ER model and discuss how its features allow us to model a wide range of data faithfully.

\Ve begin with an overview of database design in Section 2.1 in order to motivate our discussion of the ER model. \Vithin the larger context of the overall design process, the ER model is used in a phase called *conceptual database design.* We then introduce the ER model in Sections 2.2, 2.3, and 2.4. In Section 2.5, we discuss database design issues involving the ER model. We briefly discuss conceptual database design for large enterprises in Section 2.6. In Section 2.7, we present an overview of UML, a design and modeling approach that is more general in its scope than the ER model.

In Section 2.8, we introduce a case study that is used as a running example throughout the book. The case study is an end-to-end database design for an Internet shop. We illustrate the first two steps in database design (requirements analysis and conceptual design) in Section 2.8. In later chapters, we extend this case study to cover the remaining steps in the design process.

We note that many variations of ER diagrams are in use and no widely accepted standards prevail. The presentation in this chapter is representative of the family of ER models and includes a selection of the most popular features.

2.1 DATABASE DESIGN AND ER DIAGRAMS

We begin our discussion of database design by observing that this is typically just one part, although a central part in data-intensive applications, of a larger software system design. Our primary focus is the design of the database, however, and we will not discuss other aspects of software design in any detail. We revisit this point in Section 2.7.

The database design process can be divided into six steps. The ER model is most relevant to the first three steps.

1. Requirements Analysis: The very first step in designing a database application is to understand what data is to be stored in the database, what applications must be built on top of it, and what operations are most frequent and subject to performance requirements. In other words, we must find out what the users want from the database. This is usually an informal process that involves discussions with user groups, a study of the current operating environment and how it is expected to change, analysis of any available documentation on existing applications that are expected to be replaced or complemented by the database, and so on.

Database Design Tools: Design tools are available from RDBMS vendors as well as third-party vendors. For example, see the following link for details on design and analysis tools from Sybase:
http://www.sybase.com/products/application_tools
The following provides details on Oracle's tools:
http://www.oracle.com/tools

Several methodologies have been proposed for organizing and presenting the information gathered in this step, and some automated tools have been developed to support this process.

2. Conceptual Database Design: The information gathered in the requirements analysis step is used to develop a high-level description of the data to be stored in the database, along with the constraints known to hold over this data. This step is often carried out using the ER model and is discussed in the rest of this chapter. The ER model is one of several high-level, or semantic, data models used in database design. The goal is to create a simple description of the data that closely matches how users and developers think of the data (and the people and processes to be represented in the data). This facilitates discussion among all the people involved in the design process, even those who have no technical background. At the same time, the initial design must be sufficiently precise to enable a straightforward translation into a data model supported by a commercial database system (which, in practice, means the relational model).

3. Logical Database Design: We must choose a DBMS to implement our database design, and convert the conceptual database design into a database schema in the data model of the chosen DBMS. We will consider only relational DBMSs, and therefore, the task in the logical design step is to convert an ER schema into a relational database schema. We discuss this step in detail in Chapter 3; the result is a conceptual schema, sometimes called the logical schema, in the relational data model.

2.1.1 Beyond ER Design

The ER diagram is just an approximate description of the data, constructed through a subjective evaluation of the information collected during requirements analysis. A more careful analysis can often refine the logical schema obtained at the end of Step 3. Once we have a good logical schema, we must consider performance criteria and design the physical schema. Finally, we must address security issues and ensure that users are able to access the data they need, but not data that we wish to hide from them. The remaining three steps of database design are briefly described next:

4. Schema Refinement: The fourth step ill database design is to analyze the collection of relations in our relational database schema to identify potential problems, and to refine it. In contrast to the requirements analysis and conceptual design steps, which are essentially subjective, schema refinement can be guided by some elegant and powerful theory. We discuss the theory of *normalizing* relations-restructuring them to ensure some desirable properties-in Chapter 19.

5. Physical Database Design: In this step, we consider typical expected workloads that our database must support and further refine the database design to ensure that it meets desired performance criteria. This step may simply involve building indexes on some tables and clustering some tables, or it may involve a substantial redesign of parts of the database schema obtained from the earlier design steps. We discuss physical design and database tuning in Chapter 20.

6. Application and Security Design: Any software project that involves a DBMS must consider aspects of the application that go beyond the database itself. Design methodologies like UML (Section 2.7) try to address the complete software design and development cycle. Briefly, we must identify the entities (e.g., users, user groups, departments) and processes involved in the application. We must describe the role of each entity in every process that is reflected in some application task, as part of a complete workflow for that task. For each role, we must identify the parts of the database that must be accessible and the parts of the database that must *not* be accessible, and we must take steps to ensure that these access rules are enforced. A DBMS provides several mechanisms to assist in this step, and we discuss this in Chapter 21.

In the implementation phase, we must code each task in an application language (e.g., Java), using the DBlVIS to access data. We discuss application development in Chapters 6 and 7.

In general, our division of the design process into steps should be seen as a classification of the *kinds* of steps involved in design. Realistically, although we might begin with the six step process outlined here, a complete database design will probably require a subsequent tuning phase in which all six kinds of design steps are interleaved and repeated until the design is satisfactory.

2.2 ENTITIES, ATTRIBUTES, AND ENTITY SETS

An entity is an object in the real world that is distinguishable frQm other objects. Examples include the following: the Green Dragonzord toy, the toy department, the manager of the toy department, the home address of the rnan-

agel' of the toy department. It is often useful to identify a collection of similar entities. Such a collection is called an entity set. Note that entity sets need not be disjoint; the collection of toy department employees and the collection of appliance department employees may both contain employee John Doe (who happens to work in both departments). We could also define an entity set called Employees that contains both the toy and appliance department employee sets.

An entity is described using a set of attributes. All entities in a given entity set have the same attributes; this is what we mean by *similar.* (This statement is an oversimplification, as we will see when we discuss inheritance hierarchies in Section 2.4.4, but it suffices for now and highlights the main idea.) Our choice of attributes reflects the level of detail at which we wish to represent information about entities. For example, the Employees entity set could use name, social security number (ssn), and parking lot (lot) as attributes. In this case we will store the name, social security number, and lot number for each employee. However, we will not store, say, an employee's address (or gender or age).

For each attribute associated with an entity set, we must identify a domain of possible values. For example, the domain associated with the attribute *name* of Employees might be the set of 20-character strings.[1] As another example, if the company rates employees on a scale of 1 to 10 and stores ratings in a field called *mting,* the associated domain consists of integers 1 through 10. Further, for each entity set, we choose a *key.* A key is a minimal set of attributes whose values uniquely identify an entity in the set. There could be more than one candidate key; if so, we designate one of them as the primary key. For now we assume that each entity set contains at least one set of attributes that uniquely identifies an entity in the entity set; that is, the set of attributes contains a key. We revisit this point in Section 2.4.3.

The Employees entity set with attributes *ssn, name,* and *lot* is shown in Figure 2.1. An entity set is represented by a rectangle, and an attribute is represented by an oval. Each attribute in the primary key is underlined. The domain information could be listed along with the attribute name, but we omit this to keep the figures compact. The key is *ssn.*

2.3 RELATIONSHIPS AND RELATIONSHIP SETS

A relationship is an association among two or more entities. For example, we may have the relationship that Attishoo works in the pharmacy department.

iTo avoid confusion, we assume that attribute names do not repeat across entity sets. This is not a real limitation because we can always use the entity set name to resolve ambiguities if the same attribute name is used in more than one entity set.

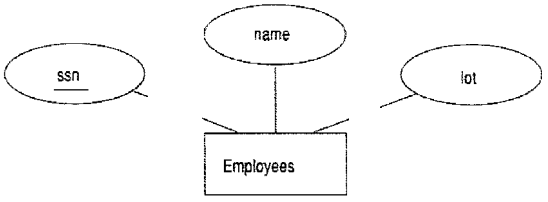


Figure 2.1 The Employees Entity Set

As with entities, we may wish to collect a set of similar relationships into a **relationship** set. A relationship set can be thought of as a set of n-tuples:

$$\{(e_1, \ldots, e_n) \mid e_1 \in E_1, \ldots, e_n \in E_n\}$$

Each n-tuple denotes a relationship involving *n* entities e_1 through *en,* where entity e_i is in entity set E_i. In Figure 2.2 we show the relationship set Works_In, in which each relationship indicates a department in which an employee works. Note that several relationship sets might involve the same entity sets. For example, we could also have a Manages relationship set involving Employees and Departments.

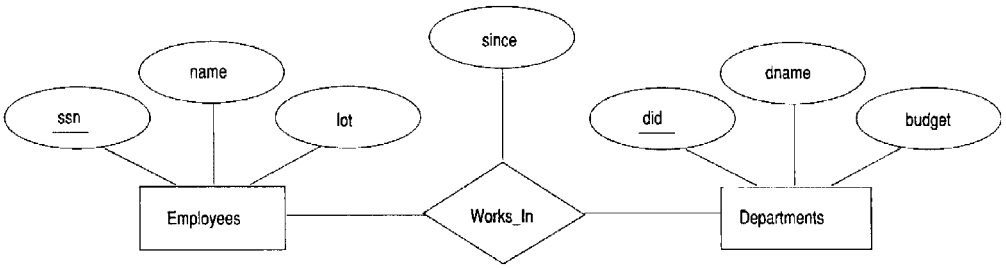


Figure 2.2 The Works_In Relationship Set

A relationship can also have **descriptive attributes.** Descriptive attributes are used to record information about the relationship, rather than about any one of the participating entities; for example, we may wish to record that Attishoo works in the pharmacy department as of January 1991. This information is captured in Figure 2.2 by adding an attribute, *since,* to Works_In. A relationship must be uniquely identified by the participating entities, without reference to the descriptive attributes. In the Works_In relationship set, for example, each Works_In relationship must be uniquely identified by the combination of employee *ssn* and department *d'id.* Thus, for a given employee-department pair, we cannot have more than one associated *since* value.

An **instance** of a relationship set is a set of relationships. Intuitively, an instance can be thought of as a 'snapshot' of the relationship set at some instant

in time. An instance of the Works_In relationship set is shown in Figure 2.3. Each Employees entity is denoted by its *ssn*, and each Departments entity is denoted by its *did*, for simplicity. The *since* value is shown beside each relationship. (The 'many-te-many' and 'total participation' comments in the figure are discussed later, when we discuss integrity constraints.)

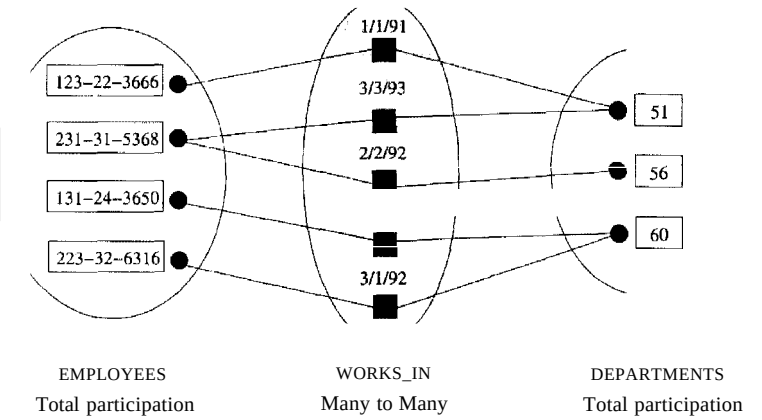


Figure 2.3 An Instance of the Works_In Relationship Set

As another example of an ER diagram, suppose that each department has offices in several locations and we want to record the locations at which each employee works. This relationship is **ternary** because we must record an association between an employee, a department, and a location. The ER diagram for this variant of Works_In, which we call Works.ln2, is shown in Figure 2.4.

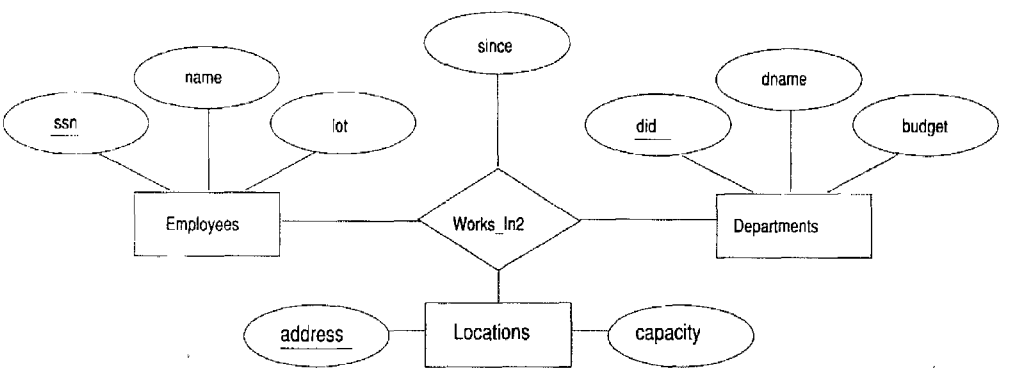


Figure 2.4 A Ternary Relationship Set

The entity sets that participate in a relationship set need not be distinct; sometimes a relationship might involve two entities in the same entity set. For example, consider the Reports_To relationship set shown in Figure 2.5. Since

employees report to other employees, every relationship in Reports_To is of the form $(emp_1, emp2)$, where both *empI* and *empz* are entities in Employees. However, they play different roles: *ernpI* reports to the managing employee *emp2,* which is reflected in the role indicators *supervisor* and *subordinate* in Figure 2.5. If an entity set plays more than one role, the role indicator concatenated with an attribute name from the entity set gives us a unique name for each attribute in the relationship set. For example, the Reports_To relationship set has attributes corresponding to the *ssn* of the supervisor and the *ssn* of the subordinate, and the names of these attributes are *supervisor_ssn* and *subordinate_ssn.*

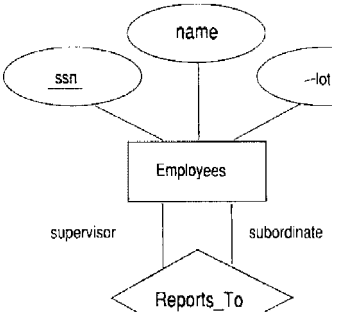


Figure 2.5 The Reports_To Relationship Set

2.4 ADDITIONAL FEATURES OF THE ER MODEL

We now look at some of the constructs in the ER model that allow us to describe some subtle properties of the data. The expressiveness of the ER model is a big reason for its widespread lise.

2.4.1 Key Constraints

Consider the Works_In relationship shown in Figure 2.2. An employee can work in several departments, and a department can have several employees, as illustrated in the vVorks_In instance shown in Figure 2.3. Employee 231-31-5368 has worked in Department 51 since 3/3/93 and in Department 56 since 2/2/92. Department 51 has two employees.

Now consider another relationship set called Manages between the Employees and Departments entity sets such that each department has at most one manager, although a single employee is allowed to manage more than one department. The restriction that each department has at most one manager is

an example of a key constraint, and it implies that each Departments entity appears in at most one Manages relationship in any allowable instance of Manages. This restriction is indicated in the ER diagram of Figure 2.6 by using an arrow from Departments to Manages. Intuitively, the arrow states that given a Departments entity, we can uniquely determine the Manages relationship in which it appears.

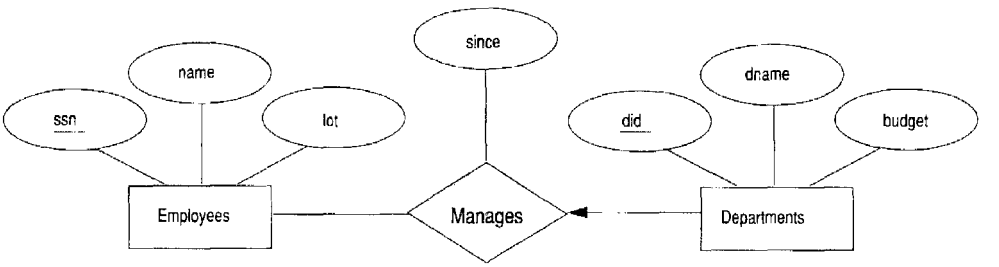


Figure 2.6 Key Constraint on Manages

An instance of the Manages relationship set is shown in Figure 2.7. While this is also a potential instance for the Works_In relationship set, the instance of Works_In shown in Figure 2.3 violates the key constraint on Manages.

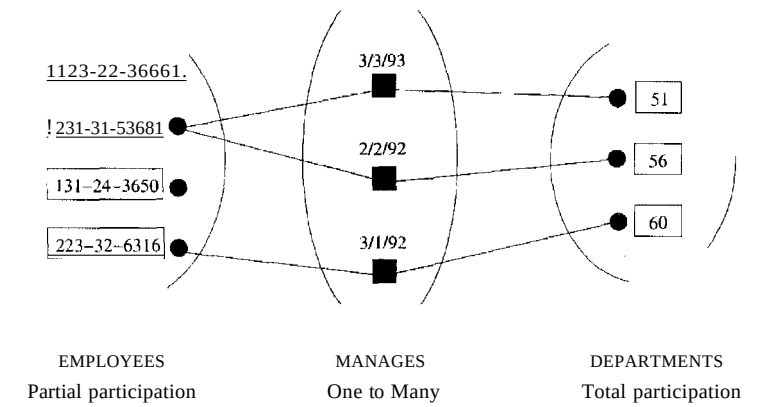


Figure 2.7 An Instance of the Manages Relationship Set

A relationship set like Manages is sometimes said to be one-to-many, to indicate that *one* employee can be associated with *many* departments (in the capacity of a manager), whereas each department can be associated with at most one employee as its manager. In contrast, the Works_In relationship set, in which an employee is allowed to work in several departments and a department is allowed to have several employees, is said to be many-to-many.

If we add the restriction that each employee can manage at most one department to the Manages relationship set, which would be indicated by adding an arrow from Employees to lVlanages in Figure 2.6, we have a one-to-one relationship set.

Key Constraints for Ternary Relationships

We can extend this convention-and the underlying key constraint concept-to relationship sets involving three or more entity sets: If an entity set E has a key constraint in a relationship set R, each entity in an instance of E appears in at most one relationship in (a corresponding instance of) R. To indicate a key constraint on entity set E in relationship set R, we draw an arrow from E to R.

In Figure 2.8, we show a ternary relationship with key constraints. Each employee works in at most one department and at a single location. An instance of the Works_In3 relationship set is shown in Figure 2.9. Note that each department can be associated with several employees and locations and each location can be associated with several departments and employees; however, each employee is associated with a single department and location.

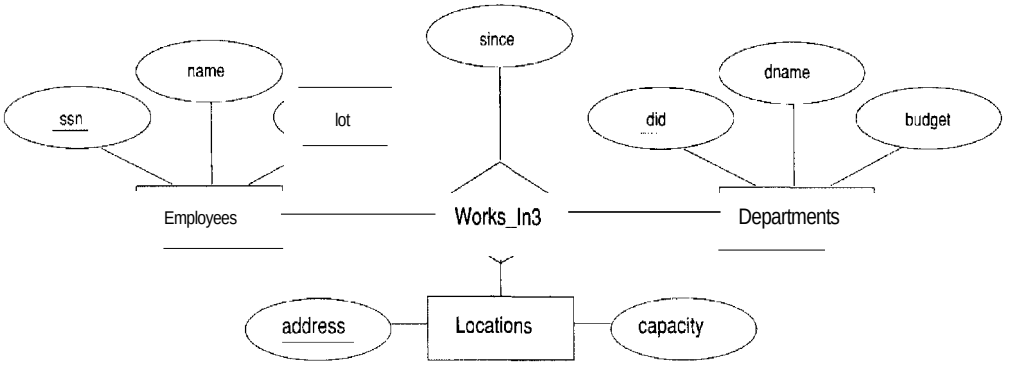


Figure 2.8 A Ternary Relationship Set with Key Constraints

2.4.2 Participation Constraints

The key constraint on Manages tells us that a department has at most one manager. A natural question to ask is whether every department has a Inanagel'. Let us say that every department is required to have a manager. This requirement is an example of a participation constraint; the participation of the entity set Departments in the relationship set Manages is said to be total. A participation that is not total is said to be partial. As an example, the

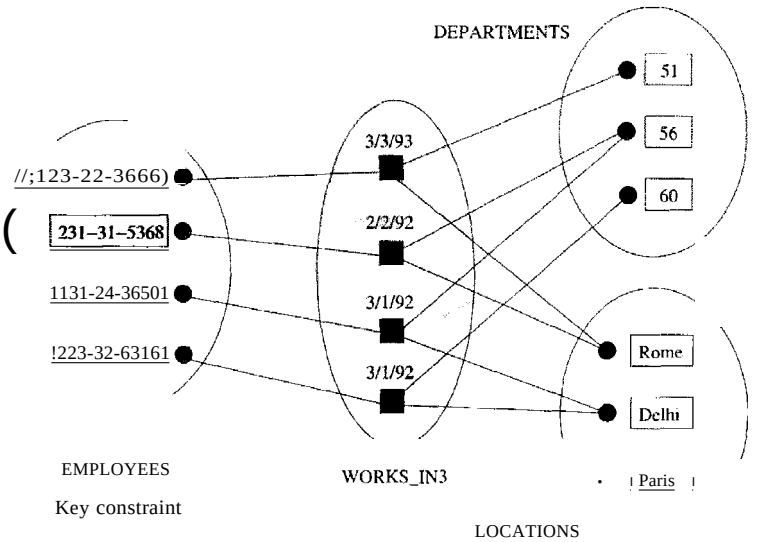


Figure 2.9 An Instance of Works_In3

participation of the entity set Employees in Manages is partial, since not every employee gets to manage a department.

Revisiting the Works..ln relationship set, it is natural to expect that each employee works in at least one department and that each department has at least one employee. This means that the participation of both Employees and Departments in Works..ln is total. The ER diagram in Figure 2.10 shows both the Manages and Works..ln relationship sets and all the given constraints. If the participation of an entity set in a relationship set is total, the two are connected by a thick line; independently, the presence of an arrow indicates a key constraint. The instances of Works_In and Manages shown in Figures 2.3 and 2.7 satisfy all the constraints in Figure 2.10.

2.4.3 Weak Entities

Thus far, we have assumed that the attributes associated with an entity set include a key. This assumption does not always hold. For example, suppose that employees can purchase insurance policies to cover their dependents. We wish to record information about policies, including who is covered by each policy, but this information is really our only interest in the dependents of an employee. If an employee quits, any policy owned by the employee is terminated and we want to delete all the relevant policy and dependent information from the database.

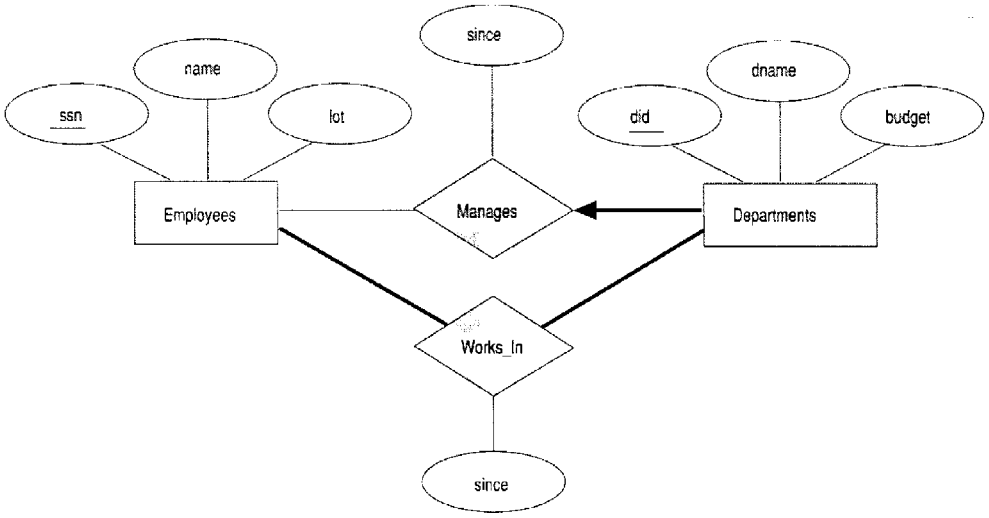


Figure 2.10 Manages and Works_In

We might choose to identify a dependent by name alone in this situation, since it is reasonable to expect that the dependents of a given employee have different names. Thus the attributes of the Dependents entity set might be *pname* and *age*. The attribute *pname* does *not* identify a dependent uniquely. Recall that the key for Employees is *ssn;* thus we might have two employees called Smethurst and each might have a son called Joe.

Dependents is an example of a weak entity set. A weak entity can be identified uniquely only by considering some of its attributes in conjunction with the primary key of another entity, which is called the identifying owner.

The following restrictions must hold:

- The owner entity set and the weak entity set must participate in a one-to-many relationship set (one owner entity is associated with one or more weak entities, but each weak entity has a single owner). This relationship set is called the identifying relationship set of the weak entity set.
- The weak entity set must have total participation in the identifying relationship set.

For example, a Dependents entity can be identified uniquely only if we take the key of the *owning* Employees entity and the *pname* of the Dependents entity. The set of attributes of a weak entity set that uniquely identify a weak entity for a given owner entity is called a *partial key* of the weak entity set. In our example, *pname* is a partial key for Dependents.

The Dependents weak entity set and its relationship to Employees is shown in Figure 2.1.1. The total participation of Dependents in Policy is indicated by linking them with a dark line. The arrow from Dependents to Policy indicates that each Dependents entity appears in at most one (indeed, exactly one, because of the participation constraint) Policy relationship. To underscore the fact that Dependents is a weak entity and Policy is its identifying relationship, we draw both with dark lines. To indicate that *pname* is a partial key for Dependents, we underline it using a broken line. This means that there may well be two dependents with the same *pname* value.

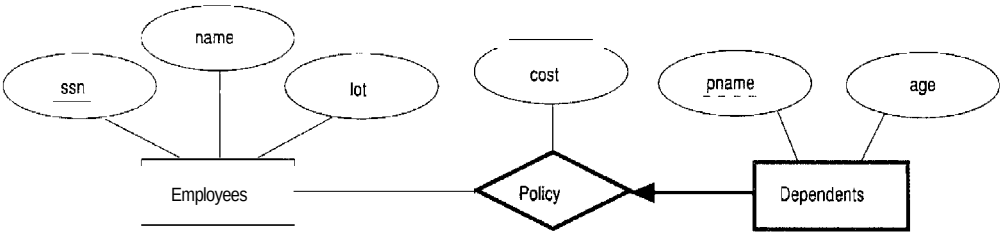


Figure 2.11 A Weak Entity Set

2.4.4 Class Hierarchies

Sometimes it is natural to classify the entities in an entity set into subclasses. For example, we might want to talk about an Hourly_Emps entity set and a ContracLEmps entity set to distinguish the basis on which they are paid. We might have attributes *hours_worked* and *hourly_wage* defined for Hourly_Emps and an attribute *contractid* defined for ContracLEmps.

We want the semantics that every entity in one of these sets is also an Employees entity and, as such, must have all the attributes of Employees defined. Therefore, the attributes defined for an Hourly_Emps entity are the attributes for Employees plus Hourly_Emps. We say that the attributes for the entity set Employees are **inherited** by the entity set Hourly_Emps and that Hourly_Emps ISA (read *is a)* Employees. In addition-and in contrast to class hierarchies in programming languages such as C++—there is a constraint on queries over instances of these entity sets: A query that asks for all Employees entities must consider all Hourly_Emps and ContracLEmps entities as well. Figure 2.12 illustrates,the class hierarchy.

The entity set Employees may also be classified using a different criterion. For example, we might identify a subset of employees as SenioLEmps. We can rnodify Figure 2.12 to reflect this change by adding a second ISA node as a child of Employees and making SenioLEmps a child of this node. Each of these entity sets might be classified further, creating a multilevel ISA hierarchy.

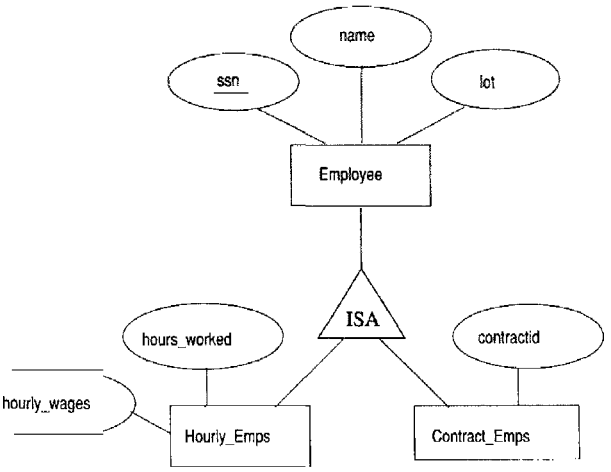


Figure 2.12 Class Hierarchy

A class hierarchy can be viewed in one of two ways:

- Employees is specialized into subclasses. Specialization is the process of identifying subsets of an entity set (the superclass) that share some distinguishing characteristic. Typically, the superclass is defined first, the subclasses are defined next, and subclass-specific attributes and relationship sets are then added.
- Hourly_Emps and ContracLEmps are generalized by Employees. As another example, two entity sets Motorboats and Cars may be generalized into an entity set Motor_Vehicles. Generalization consists of identifying some common characteristics of a collection of entity sets and creating a new entity set that contains entities possessing these common characteristics. Typically, the subclasses are defined first, the superclass is defined next, and any relationship sets that involve the superclass are then defined.

We can specify two kinds of constraints with respect to ISA hierarchies, namely, *overlap* and *covering* constraints. Overlap constraints determine whether two subclasses are allowed to contain the same entity. For example, can Attishoo be both an Hourly_Emps entity and a ContracLEmps entity? Intuitively, no. Can he be both a ContracLEmps entity and a Senior_Emps entity? Intuitively, yes. We denote this by writing 'Contract_Emps OVERLAPS Senior_Emps.' In the absence of such a statement, we assume by default that entity sets are constrained to have no overlap.

Covering constraints determine whether the entities in the subclasses collectively include all entities in the superclass. For example, does every Employees

entity have to belong to one of its subclasses? Intuitively, no. Does every Motor_Vehicles entity have to be either a Motorboats entity or a Cars entity? Intuitively, yes; a characteristic property of generalization hierarchies is that every instance of a superclass is an instance of a subclass. We denote this by writing 'Motorboats AND Cars COVER Motor-Vehicles.' In the absence of such a statement, we assume by default that there is no covering constraint; we can have motor vehicles that are not motorboats or cars.

There are two basic reasons for identifying subclasses (by specialization or generalization):

1. We might want to add descriptive attributes that make sense only for the entities in a subclass. For example, *hourly_wages* does not make sense for a ContracLEmps entity, whose pay is determined by an individual contract.
2. We might want to identify the set of entities that participate in some relationship. For example, we might wish to define the Manages relationship so that the participating entity sets are Senior_Emps and Departments, to ensure that only senior employees can be managers. As another example, Motorboats and Cars may have different descriptive attributes (say, tonnage and number of doors), but as Motor_Vehicles entities, they must be licensed. The licensing information can be captured by a Licensed_To relationship between Motor_Vehicles and an entity set called Owners.

2.4.5 Aggregation

As defined thus far, a relationship set is an association between entity sets. Sometimes, we have to model a relationship between a collection of entities and *relationships.* Suppose that we have an entity set called Projects and that each Projects entity is sponsored by one or more departments. The Sponsors relationship set captures this information. A department that sponsors a project might assign employees to monitor the sponsorship. Intuitively, Monitors should be a relationship set that associates a Sponsors relationship (rather than a Projects or Departments entity) with an Employees entity. However, we have defined relationships to associate two or more *entities.*

To define a relationship set such as Monitors, we introduce a new feature of the ER model, called *aggregation.* Aggregation allows us to indicate that a relationship set (identified through a dashed box) participates in another relationship set. This is illustrated in Figure 2.13, with a dashed box around Sponsors (and its participating entity sets) used to denote aggregation. This effectively allows us to treat Sponsors as an entity set for purposes of defining the Monitors relationship set.

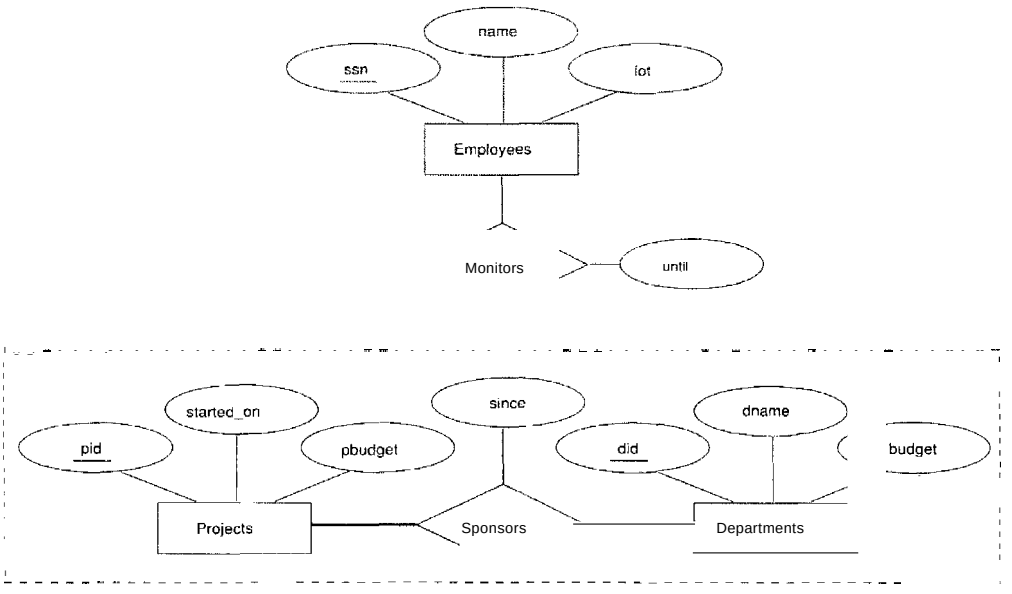


Figure 2.13 Aggregation

When should we use aggregation? Intuitively, we use it when we need to express a relationship among relationships. But can we not express relationships involving other relationships without using aggregation? In our example, why not make Sponsors a ternary relationship? The answer is that there are really two distinct relationships, Sponsors and Monitors, each possibly with attributes of its own. For instance, the Monitors relationship has an attribute *until* that records the date until when the employee is appointed as the sponsorship monitor. Compare this attribute with the attribute *since* of Sponsors, which is the date when the sponsorship took effect. The use of aggregation versus a ternary relationship may also be guided by certain integrity constraints, as explained in Section 2.5.4.

2.5 CONCEPTUAL DESIGN WITH THE ER MODEL

Developing an ER diagram presents several choices, including the following:

- Should a concept be modeled as an entity or an attribute?
- Should a concept be modeled as an entity or a relationship?
- What are the relationship sets and their participating entity sets? Should we use binary or ternary relationships?
- Should we use aggregation?

\Ve now discuss the issues involved in making these choices.

2.5.1 Entity versus Attribute

\While identifying the attributes of an entity set, it is sometimes not clear whether a property should be modeled as an attribute or as an entity set (and related to the first entity set using a relationship set). For example, consider adding address information to the Employees entity set. One option is to use an attribute *address*. This option is appropriate if we need to record only one address per employee, and it suffices to think of an address as a string. An alternative is to create an entity set called Addresses and to record associations between employees and addresses using a relationship (say, Has_Address). This more complex alternative is necessary in two situations:

- We have to record more than one address for an employee.
- We want to capture the structure of an address in our ER diagram. For example, we might break down an address into city, state, country, and Zip code, in addition to a string for street information. By representing an address as an entity with these attributes, we can support queries such as "Find all employees with an address in Madison, WI."

For another example of when to model a concept as an entity set rather than an attribute, consider the relationship set (called WorksJ:n4) shown in Figure 2.14.

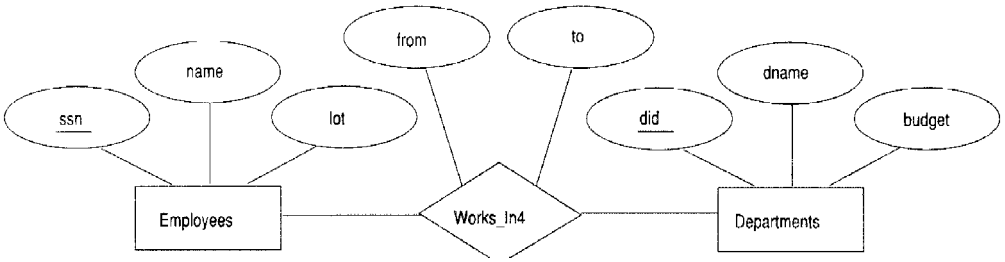


Figure 2.14 The Works_In4 Relationship Set

It differs from the \Vorks_In relationship set of Figure 2.2 only in that it has attributes *from* and *to*, instead of *since*. Intuitively, it records the interval during which an employee works for a department. Now suppose that it is possible for an employee to work in a given department over more than one period.

This possibility is ruled out by the ER diagram's semantics, because a relationship is uniquely identified by the participating entities (recall from Section

2.3). The problem is that we want to record several values for the descriptive attributes for each instance of the Works_In2 relationship. (This situation is analogous to wanting to record several addresses for each employee.) We can address this problem by introducing an entity set called, say, Duration, with attributes *from* and *to,* as shown in Figure 2.15.

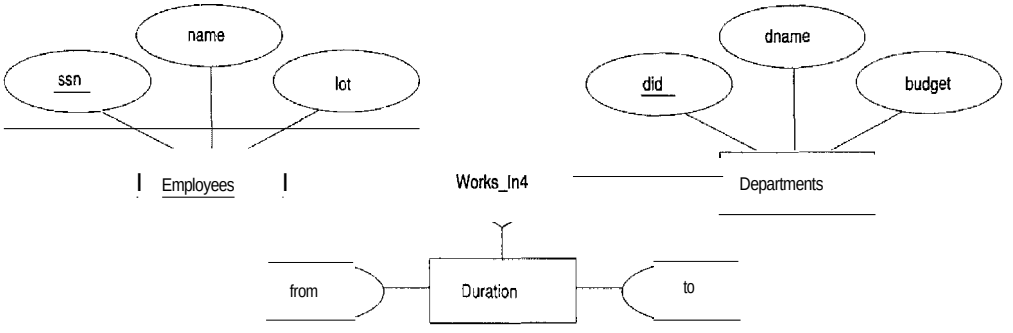


Figure 2.15 The Works_In4 Relationship Set

In some versions of the ER model, attributes are allowed to take on sets as values. Given this feature, we could make Duration an attribute of Works_In, rather than an entity set; associated with each Works_In relationship, we would have a set of intervals. This approach is perhaps more intuitive than modeling Duration as an entity set. Nonetheless, when such set-valued attributes are translated into the relational model, which does not support set-valued attributes, the resulting relational schema is very similar to what we get by regarding Duration as an entity set.

2.5.2 Entity versus Relationship

Consider the relationship set called Manages in Figure 2.6. Suppose that each department manager is given a discretionary budget *(dbudget)*, as shown in Figure 2.16, in which we have also renamed the relationship set to Manages2.

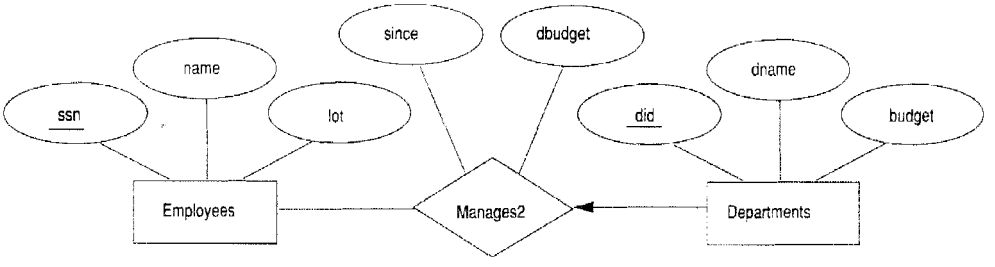


Figure 2.16 Entity versus Relationship

Given a department, we know the manager, as well as the manager's starting date and budget for that department. This approach is natural if we assume that a manager receives a separate discretionary budget for each department that he or she manages.

But what if the discretionary budget is a sum that covers *all* departments managed by that employee? In this case, each Manages2 relationship that involves a given employee will have the same value in the *dbudget* field, leading to redundant storage of the same information. Another problem with this design is that it is misleading; it suggests that the budget is associated with the relationship, when it is actually associated with the manager.

We can address these problems by introducing a new entity set called Managers (which can be placed below Employees in an ISA hierarchy, to show that every manager is also an employee). The attributes *since* and *dbudget* now describe a manager entity, as intended. As a variation, while every manager has a budget, each manager may have a different starting date (as manager) for each department. In this case *dbudget* is an attribute of Managers, but *since* is an attribute of the relationship set between managers and departments.

The imprecise nature of ER modeling can thus make it difficult to recognize underlying entities, and we might associate attributes with relationships rather than the appropriate entities. In general, such mistakes lead to redundant storage of the same information and can cause many problems. We discuss redundancy and its attendant problems in Chapter 19, and present a technique called *normalization* to eliminate redundancies from tables.

2.5.3 Binary versus Ternary Relationships

Consider the ER diagram shown in Figure 2.17. It models a situation in which an employee can own several policies, each policy can be owned by several employees, and each dependent can be covered by several policies.

Suppose that we have the following additional requirements:

- A policy cannot be owned jointly by two or more employees.
- Every policy must be owned by some employee.
- Dependents is a weak entity set, and each dependent entity is uniquely identified by taking *pname* in conjunction with the *policyid* of a policy entity (which, intuitively, covers the given dependent).

The first requirement suggests that we impose a key constraint on Policies with respect to Covers, but this constraint has the unintended side effect that a

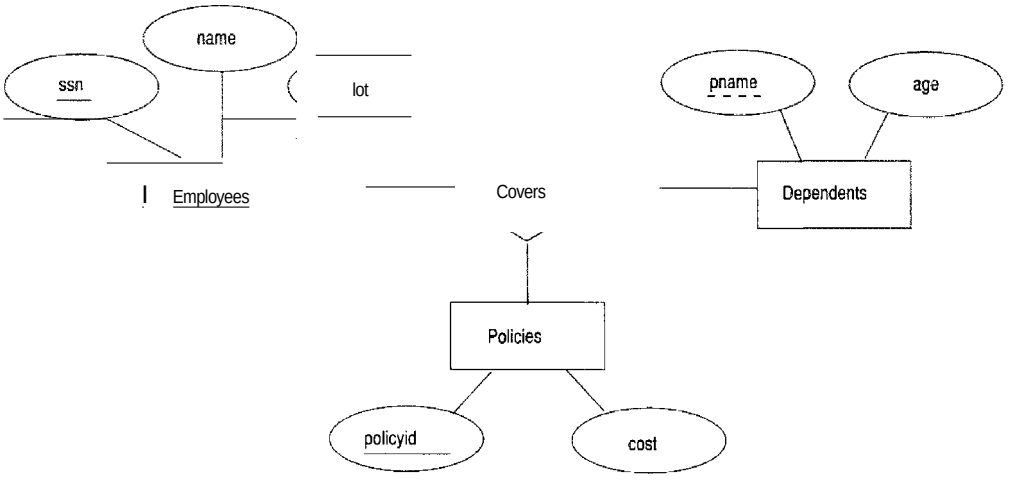


Figure 2.17 Policies as an Entity Set

policy can cover only one dependent. The second requirement suggests that we impose a total participation constraint on Policies. This solution is acceptable if each policy covers at least one dependent. The third requirement forces us to introduce an identifying relationship that is binary (in our version of ER diagrams, although there are versions in which this is not the case).

Even ignoring the third requirement, the best way to model this situation is to use two binary relationships, as shown in Figure 2.18.

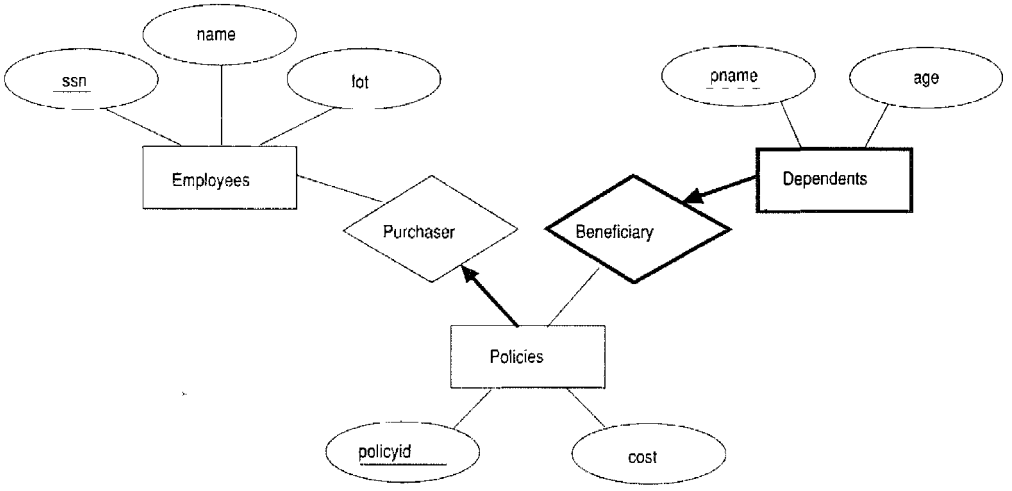


Figure 2.18 Policy Revisited

This example really has two relationships involving Policies, and our attempt to use a single ternary relationship (Figure 2.17) is inappropriate. There are situations, however, where a relationship inherently associates more than two entities. We have seen such an example in Figures 2.4 and 2.15.

As a typical example of a ternary relationship, consider entity sets Parts, Suppliers, and Departments, and a relationship set Contracts (with descriptive attribute *qty)* that involves all of them. A contract specifies that a supplier will supply (some quantity of) a part to a department. This relationship cannot be adequately captured by a collection of binary relationships (without the use of aggregation). With binary relationships, we can denote that a supplier 'can supply' certain parts, that a department 'needs' some parts, or that a department 'deals with' a certain supplier. No combination of these relationships expresses the meaning of a contract adequately, for at least two reasons:

- The facts that supplier S can supply part P, that department D needs part P, and that D will buy from S do not necessarily imply that department D indeed buys part P from supplier S!
- We cannot represent the *qty* attribute of a contract cleanly.

2.5.4 Aggregation versus Ternary Relationships

As we noted in Section 2.4.5, the choice between using aggregation or a ternary relationship is mainly determined by the existence of a relationship that relates a *relationship set* to an entity set (or second relationship set). The choice may also be guided by certain integrity constraints that we want to express. For example, consider the ER diagram shown in Figure 2.13. According to this diagram, a project can be sponsored by any number of departments, a department can sponsor one or more projects, and each sponsorship is monitored by one or more employees. If we don't need to record the *until* attribute of Monitors, then we might reasonably use a ternal'Y relationship, say, Sponsors2, as shown in Figure 2.19.

Consider the constraint that each sponsorship (of a project by a department) be monitored by at most one employee. We cannot express this constraint in terms of the Sponsors2 relationship set. On the other hand, we can easily express the cOnstraint by drawing an arrow from the aggregated relationship Sponsors to the relationship Monitors in Figure 2.13. Thus, the presence of such a constraint serves as another reason for using aggregation rather than a ternary relationship set.

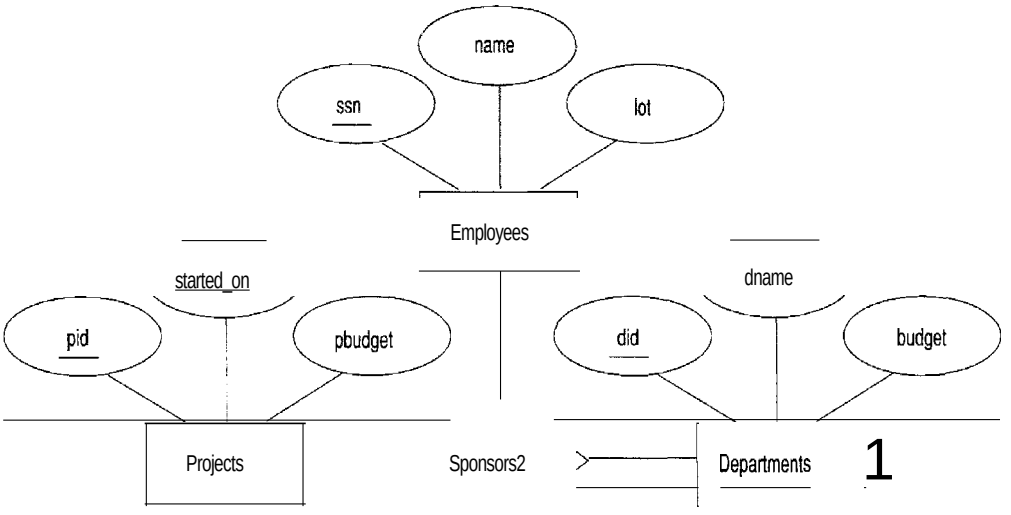


Figure 2.19 Using a Ternary Relationship instead of Aggregation

2.6 CONCEPTUAL DESIGN FOR LARGE ENTERPRISES

We have thus far concentrated on the constructs available in the ER model for describing various application concepts and relationships. The process of conceptual design consists of more than just describing small fragments of the application in terms of ER diagrams. For a large enterprise, the design may require the efforts of more than one designer and span data and application code used by a number of user groups. Using a high-level, semantic data model, such as ER diagrams, for conceptual design in such an environment offers the additional advantage that the high-level design can be diagrammatically represented and easily understood by the many people who must provide input to the design process.

An important aspect of the design process is the methodology used to structure the development of the overall design and ensure that the design takes into account all user requirements and is consistent. The usual approach is that the requirements of various user groups are considered, any conflicting requirements are somehow resolved, and a single set of global requirements is generated at the end of the.requirements analysis phase. Generating a single set of global requirements is a difficult task, but it allows the conceptual design phase to proceed with the development of a logical schema that spans all the data and applications throughout the enterprise.

An alternative approach is to develop separate conceptual schemas for different user groups and then *integrate* these conceptual schemas. To integrate multi-

pIe conceptual schemas, we must establish correspondences between entities, relationships, and attributes, and we must resolve numerous kinds of conflicts (e.g., naming conflicts, domain mismatches, differences in measurement units). This task is difficult in its own right. In some situations, schema integration cannot be avoided; for example, when one organization merges with another, existing databases may have to be integrated. Schema integration is also increasing in importance as users demand access to *heterogeneous* data sources, often maintained by different organizations.

2.7 THE UNIFIED MODELING LANGUAGE

There are many approaches to end-to-end software system design, covering all the steps from identifying the business requirements to the final specifications for a complete application, including workflow, user interfaces, and many aspects of software systems that go well beyond databases and the data stored in them. In this section, we briefly discuss an approach that is becoming popular, called the unified modeling language (UML) approach.

UML, like the ER model, has the attractive feature that its constructs can be drawn as diagrams. It encompasses a broader spectrum of the software design process than the ER model:

- Business Modeling: In this phase, the goal is to describe the business processes involved in the software application being developed.
- System Modeling: The understanding of business processes is used to identify the requirements for the software application. One part of the requirements is the database requirements.
- Conceptual Database Modeling: This step corresponds to the creation of the ER design for the database. For this purpose, UML provides many constructs that parallel the ER constructs.
- Physical Database Modeling: UML also provides pictorial representations for physical database design choices, such as the creation of table spaces and indexes. (We discuss physical database design in later chapters, but not the corresponding UML constructs.)
- Hardware System Modeling: UML diagrams can be used to describe the hardware configuration used for the application.

There are many kinds of diagrams in UML. Use case diagrams describe the actions performed by the system in response to user requests, and the people involved in these actions. These diagrams specify the external functionality that the system is expected to support.

Activity diagrams show the flow of actions in a business process. **Statechart** diagrams describe dynamic interactions between system objects. These diagrams, used in business and system modeling, describe how the external functionality is to be implemented, consistent with the business rules and processes of the enterprise.

Class diagrams are similar to ER diagrams, although they are more general in that they are intended to model *application* entities (intuitively, important program components) and their logical relationships in addition to data entities and their relationships.

Both entity sets and relationship sets can be represented as classes in UML, together with key constraints, weak entities, and class hierarchies. The term *relationship* is used slightly differently in UML, and UML's relationships are binary. This sometimes leads to confusion over whether relationship sets in an ER diagram involving three or more entity sets can be directly represented in UML. The confusion disappears once we understand that all relationship sets (in the ER sense) are represented as classes in UML; the binary UML 'relationships' are essentially just the links shown in ER diagrams between entity sets and relationship sets.

Relationship sets with key constraints are usually omitted from UML diagrams, and the relationship is indicated by directly linking the entity sets involved. For example, consider Figure 2.6. A UML representation of this ER diagram would have a class for Employees, a class for Departments, and the relationship Manages is shown by linking these two classes. The link can be labeled with a name and cardinality information to show that a department can have only one manager.

As we will see in Chapter 3, ER diagrams are translated into the relational model by mapping each entity set into a table and each relationship set into a table. Further, as we will see in Section 3.5.3, the table corresponding to a one-to-many relationship set is typically omitted by including some additional information about the relationship in the table for one of the entity sets involved. Thus, UML class diagrams correspond closely to the tables created by mapping an ER diagram.

Indeed, every class in a UML class diagram is mapped into a table in the corresponding UML database diagram. UML's **database diagrams** show how classes are represented in the database and contain additional details about the structure of the database such as integrity constraints and indexes. Links (UML's 'relationships') between UML classes lead to various integrity constraints between the corresponding tables. Many details specific to the relational model (e.g., *views, fOTe'ign keys, null-allowed fields)* and that reflect

physical design choices (e.g., indexed fields) can be modeled ill UML database diagrams.

UML's **component** diagrams describe storage aspects of the database, such as *tablespaces* and *database pa,titions)*, as well as interfaces to applications that access the database. Finally, **deployment** diagrams show the hardware aspects of the system.

Our objective in this book is to concentrate on the data stored in a database and the related design issues. To this end, we deliberately take a simplified view of the other steps involved in software design and development. Beyond the specific discussion of UML, the material in this section is intended to place the design issues that we cover within the context of the larger software design process. We hope that this will assist readers interested in a more comprehensive discussion of software design to complement our discussion by referring to other material on their preferred approach to overall system design.

2.8 CASE STUDY: THE INTERNET SHOP

We now introduce an illustrative, 'cradle-to-grave' design case study that we use as a running example throughout this book. DBDudes Inc., a well-known database consulting firm, has been called in to help Barns and Nobble (B&N) with its database design and implementation. B&N is a large bookstore specializing in books on horse racing, and it has decided to go online. DBDudes first verifies that B&N is willing and able to pay its steep fees and then schedules a lunch meeting--billed to B&N, naturally—to do requirements analysis.

2.8.1 Requirements Analysis

The owner of B&N, unlike many people who need a database, has thought extensively about what he wants and offers a concise summary:

"I would like my customers to be able to browse my catalog of books and place orders over the Internet. Currently, I take orders over the phone. I have mostly corporate customers who call me and give me the ISBN number of a book and a quantity; they often pay by credit card. I then prepare a shipment that contains the books they ordered. If I don't have enough copies in stock, I order additional copies and delay the shipment until the new copies arrive; I want to ship a customer's entire order together. My catalog includes all the books I sell. For each book, the catalog contains its ISBN number, title, author, purchase price, sales price, and the year the book was published. Most of my customers are regulars, and I have records with their names and addresses.

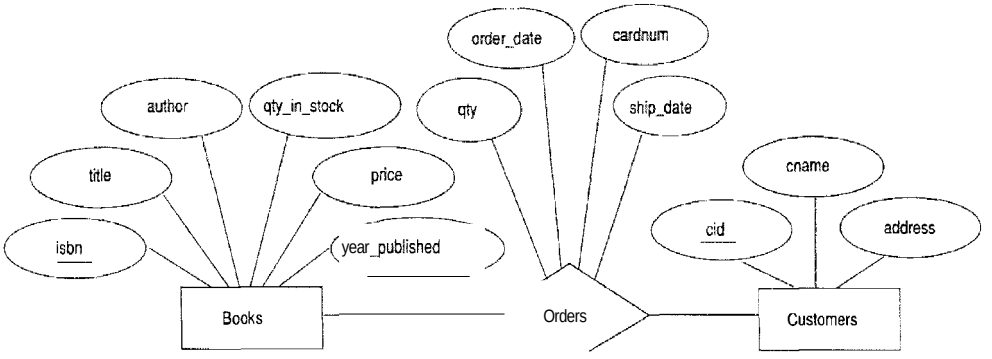


Figure 2.20 ER Diagram of the Initial Design

New customers have to call me first and establish an account before they can use my website.

On my new website, customers should first identify themselves by their unique customer identification number. Then they should be able to browse my catalog and to place orders online."

DBDudes's consultants are a little surprised by how quickly the requirements phase is completed--it usually takes weeks of discussions (and many lunches and dinners) to get this done—but return to their offices to analyze this information.

2.8.2 Conceptual Design

In the conceptual design step, DBDudes develops a high level description of the data in terms of the ER model. The initial design is shown in Figure 2.20. Books and customers are modeled as entities and related through orders that customers place. Orders is a relationship set connecting the Books and Customers entity sets. For each order, the following attributes are stored: quantity, order date, and ship date. As soon as an order is shipped, the ship date is set; until then the ship date is set to *null*, indicating that this order has not been shipped yet.

DBDudes has an internal design review at this point, and several questions are raised. To protect their identities, we will refer to the design team leader as Dude 1 and the design reviewer as Dude 2.

Dude 2: What if a customer places two orders for the same book in one day?
Dude 1: The first order is ha,ndlecl by crea.ting a new Orders relationship and

the second order is handled by updating the value of the quantity attribute in this relationship.
Dude 2: What if a customer places two orders for different books in one day?
Dude 1: No problem. Each instance of the Orders relationship set relates the customer to a different book.
Dude 2: Ah, but what if a customer places two orders for the same book on different days?
Dude 1: We can use the attribute order date of the orders relationship to distinguish the two orders.
Dude 2: Oh no you can't. The attributes of Customers and Books must jointly contain a key for Orders. So this design does not allow a customer to place orders for the same book on different days.
Dude 1: Yikes, you're right. Oh well, B&N probably won't care; we'll see.

DBDudes decides to proceed with the next phase, logical database design; we rejoin them in Section 3.8.

2.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Name the main steps in database design. What is the goal of each step? In which step is the ER model mainly used? (Section 2.1)
- Define these terms: *entity, entity set, attribute, key.* (Section 2.2)
- Define these terms: *relationship, relationship set, descriptive attributes.* (Section 2.3)
- Define the following kinds of constraints, and give an example of each: *key constraint, participation constraint.* What is a *weak entity?* What are *class hierarchies'?* What is *aggregation?* Give an example scenario motivating the use of each of these ER model design constructs. (Section 2.4)
- What guidelines would you use for each of these choices when doing ER design: \Vhether to use an attribute or an entity set, an entity or a relationship set, a binary or ternary relationship, or aggregation. (Section 2.5)
- Why is designing a database for a large enterprise especially hard? (Section 2.6)
- What is UML? How does database design fit into the overall design of a data-intensive software system? How is UML related to ER diagrams? (Section 2.7)

EXERCISES

Exercise 2.1 Explain the following terms briefly: *attribute, domain, entity, relationship, entity set, relationship set, one-to-many relationship, many-to-many relationship, participation constmint. overlap constraint, covering constraint, weak entity set, aggregation,* and *role indicator.*

Exercise 2.2 A university database contains information about professors (identified by social security number, or SSN) and courses (identified by courseid). Professors teach courses; each of the following situations concerns the Teaches relationship set. For each situation, draw an ER diagram that describes it (assuming no further constraints hold).

1. Professors can teach the same course in several semesters, and each offering must be recorded.
2. Professors can teach the same course in several semesters, and only the most recent such offering needs to be recorded. (Assume this condition applies in all subsequent questions.)
3. Every professor must teach some course.
4. Every professor teaches exactly one course (no more, no less).
5. Every professor teaches exactly one course (no more, no less), and every course must be taught by some professor.
6. Now suppose that certain courses can be taught by a team of professors jointly, but it is possible that no one professor in a team can teach the course. Model this situation, introducing additional entity sets and relationship sets if necessary.

Exercise 2.3 Consider the following information about a university database:

- Professors have an SSN, a name, an age, a rank, and a research specialty.
- Projects have a project number, a sponsor name (e.g., NSF), a starting date, an ending date, and a budget.
- Graduate students have an SSN, a name, an age, and a degree program (e.g., M.S. or Ph.D.).
- Each project is managed by one professor (known as the project's principal investigator).
- Each project is worked on by one or more professors (known as the project's co-investigators).
- Professors can manage and/or work on multiple projects.
- Each project is worked on by one or more graduate students (known as the project's research assistants).
- When graduate students work on a project, a professor must supervise their work on the project. Graduate students can work on multiple projects, in which case they will have a (potentially different) supervisor for each one.
- Departments have a department number, a department name, and a main office.
- Departments have a professor (known as the chairman) who runs the department.
- Professors work in one or more departments, and for each department that they work in, a time percentage is associated with their job.
- Graduate students have one major department in which they are working on their degree.

- Each graduate student has another, more senior graduate student (known as a student advisor) who advises him or her on what courses to take.

Design and draw an ER diagram that captures the information about the university. Use only the basic ER model here; that is, entities, relationships, and attributes. Be sure to indicate any key and participation constraints.

Exercise 2.4 A company database needs to store information about employees (identified by *ssn*, with *salary* and *phone* as attributes), departments (identified by *dna*, with *dname* and *budget* as attributes), and children of employees (with *name* and *age* as attributes). Employees *work* in departments; each department is *managed by* an employee; a child must be identified uniquely by *name* when the parent (who is an employee; assume that only one parent works for the company) is known. We are not interested in information about a child once the parent leaves the company.

Draw an ER diagram that captures this information.

Exercise 2.5 Notown Records has decided to store information about musicians who perform on its albums (as well as other company data) in a database. The company has wisely chosen to hire you as a database designer (at your usual consulting fee of $2500/day).

- Each musician that records at Notown has an SSN, a name, an address, and a phone number. Poorly paid musicians often share the same address, and no address has more than one phone.
- Each instrument used in songs recorded at Notown has a name (e.g., guitar, synthesizer, flute) and a musical key (e.g., C, B-flat, E-flat).
- Each album recorded on the Notown label has a title, a copyright date, a format (e.g., CD or MC), and an album identifier.
- Each song recorded at Notown has a title and an author.
- Each musician may play several instruments, and a given instrument may be played by several musicians.
- Each album has a number of songs on it, but no song may appear on more than one album.
- Each song is performed by one or more musicians, and a musician may perform a number of songs.
- Each album has exactly one musician who acts as its producer. A musician may produce several albums, of course.

Design a conceptual schema for Notown and draw an ER diagram for your schema. The preceding information describes the situation that the Notown database must model. Be sure to indicate all key and cardinality constraints and any assumptions you make. Identify any constraints you are unable to capture in the ER diagram and briefly explain why you could not express them.

Exercise 2.6 Computer Sciences Department frequent fliers have been complaining to Dane County Airport officials about the poor organization at the airport. As a result, the officials decided that all information related to the airport should be organized using a DBMS, and you have been hired to design the database. Your first task is to organize the information about all the airplanes stationed and maintained at the airport. The relevant information is as follows:

- Every airplane has a registration number, and each airplane is of a specific model.
- The airport accommodates a number of airplane models, and each model is identified by a model number (e.g., DC-10) and has a capacity and a weight.
- A number of technicians work at the airport. You need to store the name, SSN, address, phone number, and salary of each technician.
- Each technician is an expert on one or more plane model(s), and his or her expertise may overlap with that of other technicians. This information about technicians must also be recorded.
- Traffic controllers must have an annual medical examination. For each traffic controller, you must store the date of the most recent exam.
- All airport employees (including technicians) belong to a union. You must store the union membership number of each employee. You can assume that each employee is uniquely identified by a social security number.
- The airport has a number of tests that are used periodically to ensure that airplanes are still airworthy. Each test has a Federal Aviation Administration (FAA) test number, a name, and a maximum possible score.
- The FAA requires the airport to keep track of each time a given airplane is tested by a given technician using a given test. For each testing event, the information needed is the date, the number of hours the technician spent doing the test, and the score the airplane received on the test.

1. Draw an ER diagram for the airport database. Be sure to indicate the various attributes of each entity and relationship set; also specify the key and participation constraints for each relationship set. Specify any necessary overlap and covering constraints as well (in English).
2. The FAA passes a regulation that tests on a plane must be conducted by a technician who is an expert on that model. How would you express this constraint in the ER diagram? If you cannot express it, explain briefly.

Exercise 2.7 The Prescriptions-R-X chain of pharmacies has offered to give you a free lifetime supply of medicine if you design its database. Given the rising cost of health care, you agree. Here's the information that you gather:

- Patients are identified by an SSN, and their names, addresses, and ages must be recorded.
- Doctors are identified by an SSN. For each doctor, the name, specialty, and years of experience must be recorded.
- Each pharmaceutical company is identified by name and has a phone number.
- For each drug, the trade name and formula must be recorded. Each drug is sold by a given pharmaceutical company, and the trade name identifies a drug uniquely from among the products of that company. If a pharmaceutical company is deleted, you need not keep track of its products any longer.
- Each pharmacy has a name, address, and phone number.
- Every patient has a primary physician. Every doctor has at least one patient.
- Each pharmacy sells several drugs and has a price for each. A drug could be sold at several pharmacies, and the price could vary from one pharmacy to another.

- Doctors prescribe drugs for patients. A doctor could prescribe one or more drugs for several patients, and a patient could obtain prescriptions from several doctors. Each prescription has a date and a quantity associated with it. You can assume that, if a doctor prescribes the same drug for the same patient more than once, only the last such prescription needs to be stored.
- Pharmaceutical companies have long-term contracts with pharmacies. A pharmaceutical company can contract with several pharmacies, and a pharmacy can contract with several pharmaceutical companies. For each contract, you have to store a start date, an end date, and the text of the contract.
- Pharmacies appoint a supervisor for each contract. There must always be a supervisor for each contract, but the contract supervisor can change over the lifetime of the contract.

1. Draw an ER diagram that captures the preceding information. Identify any constraints not captured by the ER diagram.
2. How would your design change if each drug must be sold at a fixed price by all pharmacies?
3. How would your design change if the design requirements change as follows: If a doctor prescribes the same drug for the same patient more than once, several such prescriptions may have to be stored.

Exercise 2.8 Although you always wanted to be an artist, you ended up being an expert on databases because you love to cook data and you somehow confused *database* with *data baste.* Your old love is still there, however, so you set up a database company, ArtBase, that builds a product for art galleries. The core of this product is a database with a schema that captures all the information that galleries need to maintain. Galleries keep information about artists, their names (which are unique), birthplaces, age, and style of art. For each piece of artwork, the artist, the year it was made, its unique title, its type of art (e.g., painting, lithograph, sculpture, photograph), and its price must be stored. Pieces of artwork are also classified into groups of various kinds, for example, portraits, still lifes, works by Picasso, or works of the 19th century; a given piece may belong to more than one group. Each group is identified by a name (like those just given) that describes the group. Finally, galleries keep information about customers. For each customer, galleries keep that person's unique name, address, total amount of dollars spent in the gallery (very important!), and the artists and groups of art that the customer tends to like.

Draw the ER diagram for the database.

Exercise 2.9 Answer the following questions.

- Explain the following terms briefly: *UML, use case diagrams, statechart diagrams, class diagrams, database diagrams, component diagrams,* and *deployment diagrams.*
- Explain the relationship between ER diagrams and UML.

BffiLIOGRAPHIC NOTES

Several books provide a good treatment of conceptual design; these include [63J (which also contains a survey of commercial database design tools) and [730J.

The ER model was proposed by Chen [172], and extensions have been proposed in a number of subsequent papers. Generalization and aggregation were introduced in [693]. [390, 589]

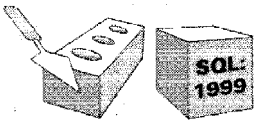
contain good surveys of semantic data models. Dynamic and temporal aspects of semantic data models are discussed in [749].

[731] discusses a design methodology based on developing an ER diagram and then translating it to the relational model. Markowitz considers referential integrity in the context of ER to relational mapping and discusses the support provided in some commercial systems (as of that date) in [513, 514].

The entity-relationship conference proceedings contain numerous papers on conceptual design, with an emphasis on the ER model; for example, [698].

The OMG home page (www.omg.org) contains the specification for UML and related modeling standards. Numerous good books discuss UML; for example [105, 278, 640] and there is a yearly conference dedicated to the advancement of UML, the International Conference on the Unified Modeling Language.

View integration is discussed in several papers, including [97, 139, 184, 244, 535, 551, 550, 685, 697, 748]. [64] is a survey of several integration approaches.



3

THE RELATIONAL MODEL

- How is data represented in the relational model?
- What integrity constraints can be expressed?
- How can data be created and modified?
- How can data be manipulated and queried?
- How can we create, modify, and query tables using SQL?
- How do we obtain a relational database design from an ER diagram?
- What are views and why are they used?
- Key concepts: relation, schema, instance, tuple, field, domain, degree, cardinality; SQL DDL, CREATE TABLE, INSERT, DELETE, UPDATE; integrity constraints, domain constraints, key constraints, PRIMARY KEY, UNIQUE, foreign key constraints, FOREIGN KEY; referential integrity maintenance, deferred and immediate constraints; relational queries; logical database design, translating ER diagrams to relations, expressing ER constraints using SQL; views, views and logical independence, security; creating views in SQL, updating views, querying views, dropping views

TABLE: An arrangement of words, numbers, or signs, or combinations of them, as in parallel columns, to exhibit a set of facts or relations in a definite, compact, and comprehensive form; a synopsis or scheme.

-----vVebster's *Dictionary of the English Language*

Codd proposed the relational data model in 1970. At that time, most database systems were based on one of two older data models (the hierarchical model

SQL. Originally developed as the query language of the pioneering System-R relational DBMS at IBM, structured query language (SQL) has become the most widely used language for creating, manipulating, and querying relational DBMSs. Since many vendors offer SQL products, there is a need for a standard that defines \official SQL.' The existence of a standard allows users to measure a given vendor's version of SQL for completeness. It also allows users to distinguish SQLfeatures specific to one product from those that are standard; an application that relies on nonstandard features is less portable.

The first SQL standard was developed in 1986 by the American National Standards Institute (ANSI) and was called SQL-86. There was a minor revision in 1989 called SQL-89 and a major revision in 1992 called SQL-92. The International Standards Organization (ISO) collaborated with ANSI to develop SQL-92. Most commercial DBMSs currently support (the core subset of) SQL-92 and are working to support the recently adopted SQL:1999 version of the standard, a major extension of SQL-92. Our coverage of SQL is based on SQL:1999, but is applicable to SQL-92 as well; features unique to SQL:1999 are explicitly noted.

and the network model); the relational model revolutionized the database field and largely supplanted these earlier models. Prototype relational database management systems were developed in pioneering research projects at IBM and DC-Berkeley by the mid-197Gs, and several vendors were offering relational database products shortly thereafter. Today, the relational model is by far the dominant data model and the foundation for the leading DBMS products, including IBM's DB2 family, Informix, Oracle, Sybase, Microsoft's Access and SQLServer, FoxBase, and Paradox. Relational database systems are ubiquitous in the marketplace and represent a multibillion dollar industry.

The relational model is very simple and elegant: a database is a collection of one or more *relations,* where each relation is a table with rows and columns. This simple tabular representation enables even novice users to understand the contents of a database, and it permits the use of simple, high-level languages to query the data. The major advantages of the relational model over the older data models are its simple data representation and the ease with which even complex queries can be expressed.

While we concentrate on the underlying concepts, we also introduce the Data Definition Language (DDL) features of SQL, the standard language for creating, manipulating, and querying data in a relational DBMS. This allows us to ground the discussion firmly in terms of real database systems.

We discuss the concept of a relation in Section 3.1 and show how to create relations using the SQL language. An important component of a data model is the set of constructs it provides for specifying conditions that must be satisfied by the data. Such conditions, called *'integrity constraints* (lGs), enable the DBIviS to reject operations that might corrupt the data. We present integrity constraints in the relational model in Section 3.2, along with a discussion of SQL support for les. We discuss how a DBMS enforces integrity constraints in Section 3.3.

In Section 3.4, we turn to the mechanism for accessing and retrieving data from the database, *query languages*, and introduce the querying features of SQL, which we examine in greater detail in a later chapter.

We then discuss converting an ER diagram into a relational database schema in Section 3.5. We introduce *views,* or tables defined using queries, in Section 3.6. Views can be used to define the external schema for a database and thus provide the support for logical data independence in the relational model. In Section 3.7, we describe SQL commands to destroy and alter tables and views.

Finally, in Section 3.8 we extend our design case study, the Internet shop introduced in Section 2.8, by showing how the ER diagram for its conceptual schema can be mapped to the relational model, and how the use of views can help in this design.

3.1 INTRODUCTION TO THE RELATIONAL MODEL

The main construct for representing data in the relational model is a relation. A relation consists of a relation schema and a relation instance. The relation instance is a table, and the relation schema describes the column heads for the table. We first describe the relation schema and then the relation instance. The schema specifies the relation's name, the name of each field (or column, or attribute), and the domain of each field. A domain is referred to in a relation schema by the domain name and has a set of associated values.

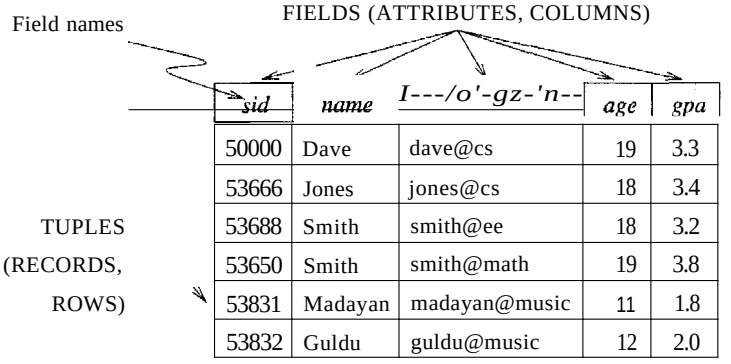
\Ve use the example of student information in a university database from Chapter 1 to illustrate the parts of a relation schema:

Students(sid: string, *name:* string, *login:* string,
age: integer, *gpa:* real)

This says, for instance, that the field named *sid* has a domain named string. The set of values associated with domain string is the set of all character strings.

We now turn to the instances of a relation. An instance of a relation is a set of tuples, also called records, in which each tuple has the same number of fields as the relation schema. A relation instance can be thought of as a *table* in which each tuple is a *row*, and all rows have the same number of fields. (The term *relation instance* is often abbreviated to just *relation,* when there is no confusion with other aspects of a relation such as its schema.)

An instance of the Students relation appears in Figure 3.1. The instance 81



sid	name	I---/o'-gz-'n--	age	gpa
50000	Dave	dave@cs	19	3.3
53666	Jones	jones@cs	18	3.4
53688	Smith	smith@ee	18	3.2
53650	Smith	smith@math	19	3.8
53831	Madayan	madayan@music	11	1.8
53832	Guldu	guldu@music	12	2.0

Figure 3.1 An Instance 81 of the Students Relation

contains six tuples and has, as we expect from the schema, five fields. Note that no two rows are identical. This is a requirement of the relational model-each relation is defined to be a *set* of unique tuples or rows.

In practice, commercial systems allow tables to have duplicate rows, but we assume that a relation is indeed a set of tuples unless otherwise noted. The order in which the rows are listed is not important. Figure 3.2 shows the same relation instance. If the fields are named, as in our schema definitions and

s'id	name	login	age	gpa
53831	Madayan	madayan@music	11	1.8
53832	Guldu	glldll@music	12	2.0
53688	Smith	smith@ee	18	3.2
53650	Smith	smith@math	19	3.8
53666	Jones	jones@cs	18	3.4
50000	Dave	dave@cs	19	3.3

Figure 3.2 An Alternative Representation of Instance 81 of Students

figures depicting relation instances, the order of fields does not matter either. However, an alternative convention is to list fields in a specific order and refer

to a field by its position. Thus, *sid* is field 1 of Students, *login* is field 3, and so on. If this convention is used, the order of fields is significant. Most database systems use a combination of these conventions. For example, in SQL, the named fields convention is used in statements that retrieve tuples and the ordered fields convention is commonly used when inserting tuples.

A relation schema specifies the domain of each field or column in the relation instance. These domain constraints in the schema specify an important condition that we want each instance of the relation to satisfy: The values that appear in a column must be drawn from the domain associated with that column. Thus, the domain of a field is essentially the *type* of that field, in programming language terms, and restricts the values that can appear in the field.

More formally, let $R(f_1{:}D1, \ldots, In{:}Dn)$ be a relation schema, and for each f_i, $1 \leq i \leq n$, let *Dami* be the set of values associated with the domain named Di. An instance of R that satisfies the domain constraints in the schema is a set of tuples with *n* fields:

$$\{ \langle f_1 : d_1, \quad ,In: d_n) \mid d_1 \in Daml' \ \ldots \ ,d_n \in Damn \}$$

The angular brackets ( ) identify the fields of a tuple. Using this notation, the first Students tuple shown in Figure 3.1 is written as *(sid:* 50000, *name:* Dave, *login:* dave@cs, *age:* 19, *gpa:* 3.3). The curly brackets {...} denote a set (of tuples, in this definition). The vertical bar | should be read 'such that,' the symbol ∈ should be read 'in,' and the expression to the right of the vertical bar is a condition that must be satisfied by the field values of each tuple in the set. Therefore, an instance of R is defined as a set of tuples. The fields of each tuple must correspond to the fields in the relation schema.

Domain constraints are so fundamental in the relational model that we henceforth consider only relation instances that satisfy them; therefore, *relation instance* means *relation instance that satisfies the domain constraints in the relation schema.*

The degree, also called arity, of a relation is the number of fields. The cardinality of a relation instance is the number of tuples in it. In Figure 3.1, the degree of the relation (the number of columns) is five, and the cardinality of this instance is six.

A relational database is a collection of relations with distinct relation names. The relational database schema is the collection of schemas for the relations in the database. 'For example, in Chapter 1, we discllssed a university database with relations called Students, Faculty, Courses, Rooms, Enrolled, Teaches, and Meets_In. An instance of a relational database is a collection of relation

instances, one per relation schema in the database schema; of course, each relation instance must satisfy the domain constraints in its schema.

3.1.1 Creating and Modifying Relations Using SQL

The SQL language standard uses the word *table* to denote *relation,* and we often follow this convention when discussing SQL. The subset of SQL that supports the creation, deletion, and modification of tables is called the Data Definition Language (DDL). Further, while there is a command that lets users define new domains, analogous to type definition commands in a programming language, we postpone a discussion of domain definition until Section 5.7. For now, we only consider domains that are built-in types, such as integer.

The CREATE TABLE statement is used to define a new table.[1] To create the Students relation, we can use the following statement:

```
CREATE TABLE Students ( sid    CHAR(20),
                        name   CHAR(30),
                        login  CHAR(20),
                        age    INTEGER,
                        gpa    REAL)
```

Tuples are inserted using the INSERT command. We can insert a single tuple into the Students table as follows:

```
INSERT
INTO    Students   (sid, name, login, age, gpa)
VALUES  (53688, 'Smith', 'smith@ee', 18, 3.2)
```

We can optionally omit the list of column names in the INTO clause and list the values in the appropriate order, but it is good style to be explicit about column names.

We can delete tuples using the DELETE command. We can delete all Students tuples with *name* equal to Smith using the command:

```
DELETE
FROM    Students S
WHERE   S.name = 'Smith'
```

[1] SQL also provides statements to destroy tables and to change the columns associated with a table; we discuss these in Section 3.7.

We can modify the column values in an existing row using the UPDATE command. For example, we can increment the age and decrement the gpa of the student with *sid 53688:*

```
UPDATE  Students S
SET     S.age = S.age + 1, S.gpa = S.gpa - 1
WHERE   S.sid = 53688
```

These examples illustrate some important points. The WHERE clause is applied first and determines which rows are to be modified. The SET clause then determines how these rows are to be modified. If the column being modified is also used to determine the new value, the value used in the expression on the right side of equals (=) is the *old* value, that is, before the modification. To illustrate these points further, consider the following variation of the previous query:

```
UPDATE  Students S
SET     S.gpa = S.gpa - 0.1
WHERE   S.gpa >= 3.3
```

If this query is applied on the instance 81 of Students shown in Figure 3.1, we obtain the instance shown in Figure 3.3.

sid	*name*	*login*	*age*	*gpa*
50000	Dave	dave@cs	19	3.2
53666	Jones	jones@cs	18	3.3
53688	Smith	smith@ee	18	3.2
53650	Smith	smith@math	19	3.7
53831	Madayan	madayan@music	11	1.8
53832	Guldu	guldu@music	12	2.0

Figure 3.3 Students Instance 81 after Update

3.2 INTEGRITY CONSTRAINTS OVER RELATIONS

A database is only as good as the information stored in it, and a DBMS must therefore help prevent the entry of incorrect information. An integrity constraint (Ic) is a condition specified on a database schema and restricts the data that can be stored in an instance of the database. If a database instance satisfies all the integrity constraints specified on the database schema, it is a legal instance. A DBMS enforces integrity constraints, in that it permits only legal instances to be stored in the database.

Integrity constraints are specified and enforced at different times:

1. When the DBA or end user defines a database schema, he or she specifies the ICs that must hold on any instance of this database.

2. When a database application is run, the DBMS checks for violations and disallows changes to the data that violate the specified ICs. (In some situations, rather than disallow the change, the DBMS might make some compensating changes to the data to ensure that the database instance satisfies all ICs. In any case, changes to the database are not allowed to create an instance that violates any IC.) It is important to specify exactly when integrity constraints are checked relative to the statement that causes the change in the data and the transaction that it is part of. We discuss this aspect in Chapter 16, after presenting the transaction concept, which we introduced in Chapter 1, in more detail.

Many kinds of integrity constraints can be specified in the relational model. We have already seen one example of an integrity constraint in the *domain constraints* associated with a relation schema (Section 3.1). In general, other kinds of constraints can be specified as well; for example, no two students have the same *sid* value. In this section we discuss the integrity constraints, other than domain constraints, that a DBA or user can specify in the relational model.

3.2.1 Key Constraints

Consider the Students relation and the constraint that no two students have the same student id. This IC is an example of a key constraint. A **key constraint** is a statement that a certain *minimal* subset of the fields of a relation is a unique identifier for a tuple. A set of fields that uniquely identifies a tuple according to a key constraint is called a **candidate key** for the relation; we often abbreviate this to just *key*. In the case of the Students relation, the (set of fields containing just the) *sid* field is a candidate key.

Let us take a closer look at the above definition of a (candidate) key. There are two parts to the definition:[2]

1. Two distinct tuples in a legal instance (an instance that satisfies all ICs, including the key constraint) cannot have identical values in all the fields of a key.

2. No subset of the set of fields in a key is a unique identifier for a tuple.

[2]The term *key* is rather overworked. In the context of access methods, we speak of *search keys*, which are quite different.

The first part of the definition means that, in *any* legal instance, the values in the key fields uniquely identify a tuple in the instance. \Vhen specifying a key constraint, the DBA or user must be sure that this constraint will not prevent them from storing a 'correct' set of tuples. (A similar comment applies to the specification of other kinds of les as well.) The notion of 'correctness' here depends on the nature of the data being stored. For example, several students may have the same name, although each student has a unique student id. If the *name* field is declared to be a key, the DBMS will not allow the Students relation to contain two tuples describing different students with the same name!

The second part of the definition means, for example, that the set of fields *{sid, name}* is not a key for Students, because this set properly contains the key *{sid}*. The set *{sid, name}* is an example of a superkey, which is a set of fields that contains a key.

Look again at the instance of the Students relation in Figure 3.1. Observe that two different rows always have different *sid* values; *sid* is a key and uniquely identifies a tuple. However, this does not hold for nonkey fields. For example, the relation contains two rows with *Smith* in the *name* field.

Note that every relation is guaranteed to have a key. Since a relation is a set of tuples, the set of *all* fields is always a superkey. If other constraints hold, some subset of the fields may form a key, but if not, the set of all fields is a key.

A relation may have several candidate keys. For example, the *login* and *age* fields of the Students relation may, taken together, also identify students uniquely. That is, *{login, age}* is also a key. It may seem that *login* is a key, since no two rows in the example instance have the same *login* value. However, the key must identify tuples uniquely in all possible legal instances of the relation. By stating that *{login, age}* is a key, the user is declaring that two students may have the same login or age, but not both.

Out of all the available candidate keys, a database designer can identify a **primary** key. Intuitively, a tuple can be referred to from elsewhere in the database by storing the values of its primary key fields. For example, we can refer to a Students tuple by storing its *sid* value. As a consequence of referring to student tuples in this manner, tuples are frequently accessed by specifying their *sid* value. In principle, we can use any key, not just the primary key, to refer to a tuple. However, using the primary key is preferable because it is what the DBMS expects this is the significance of designating a particular candidate key as a primary key and optimizes for. For example, the DBMS may create an index with the primary key fields as the search key, to make the retrieval of a tuple given its primary key value efficient. The idea of referring to a tuple is developed further in the next section.

Specifying Key Constraints in SQL

In SQL, we can declare that a subset of the columns of a table constitute a key by using the UNIQUE constraint. At most one of these candidate keys can be declared to be a *primary key*, using the PRIMARY KEY constraint. (SQL does not require that such constraints be declared for a table.)

Let us revisit our example table definition and specify key information:

```
CREATE TABLE Students ( sid      CHAR(20),
                        name     CHAR(30),
                        login    CHAR(20),
                        age      INTEGER,
                        gpa      REAL,
                        UNIQUE (name, age),
                        CONSTRAINT StudentsKey PRIMARY KEY (sid) )
```

This definition says that *sid* is the primary key and the combination of *name* and *age* is also a key. The definition of the primary key also illustrates how we can name a constraint by preceding it with CONSTRAINT *constraint-name*. If the constraint is violated, the constraint name is returned and can be used to identify the error.

3.2.2 Foreign Key Constraints

Sometimes the information stored in a relation is linked to the information stored in another relation. If one of the relations is modified, the other must be checked, and perhaps modified, to keep the data consistent. An IC involving both relations must be specified if a DBMS is to make such checks. The most common IC involving two relations is a *foreign key* constraint.

Suppose that, in addition to Students, we have a second relation:

Enrolled(studid: string, *cid:* string, *gTade:* string)

To ensure that only bona fide students can enroll in courses, any value that appears in the *studid* field of an instance of the Enrolled relation should also appear in the *sid* field of some tuple in the Students relation. The *studid* field of Enrolled is called a foreign key and refers to Students. The foreign key in the referencing relation (Enrolled, in our example) must match the primary key of the referenced relation (Students); that is, it must have the same number of columns and compatible data types, although the column names can be different.

This constraint is illustrated in Figure 3.4. As the figure shows, there may well be some Students tuples that are not referenced from Enrolled (e.g., the student with *sid=50000*). However, every *studid* value that appears in the instance of the Enrolled table appears in the primary key column of a row in the Students table.

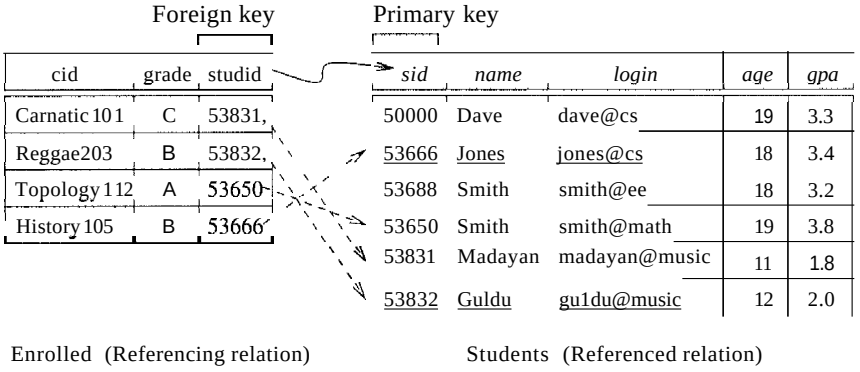


Figure 3.4 Referential Integrity

If we try to insert the tuple (55555, *Artl04, A)* into *E1,* the Ie is violated because there is no tuple in 51 with *sid* 55555; the database system should reject such an insertion. Similarly, if we delete the tuple (53666, *Jones, jones@cs, 18,* 3.4) from 51, we violate the foreign key constraint because the tuple (53666, *Historyl05, B)* in *E1* contains *studid* value 53666, the *sid* of the deleted Students tuple. The DBMS should disallow the deletion or, perhaps, also delete the Enrolled tuple that refers to the deleted Students tuple. We discuss foreign key constraints and their impact on updates in Section 3.3.

Finally, we note that a foreign key could refer to the same relation. For example, we could extend the Students relation with a column called *partner* and declare this column to be a foreign key referring to Students. Intuitively, every student could then have a partner, and the *partner* field contains the partner's *sid.* The observant reader will no doubt ask, "What if a student does not (yet) have a partnerT' This situation is handled in SQL by using a special value called null. The use of *null* in a field of a tuple rneans that value in that field is either unknown or not applicable (e.g., we do not know the partner yet or there is no partner). The appearance of *null* in a foreign key field does not violate the foreign key constraint. However, *null* values are not allowed to appear in a primary key field (because the primary key fields are used to identify a tuple uniquely). We discuss *null* values further in Chapter 5.

Specifying Foreign Key Constraints in SQL

Let us define Enrolled(studid: string, *cid:* string, *grade:* string):

```
CREATE TABLE Enrolled ( studid CHAR(20),
                        cid    CHAR(20),
                        grade  CHAR(10),
                        PRIMARY KEY (studid, cid),
                        FOREIGN KEY (studid) REFERENCES Students)
```

The foreign key constraint states that every *st'udid* value in Enrolled must also appear in Students, that is, *studid* in Enrolled is a foreign key referencing Students. Specifically, every *studid* value in Enrolled must appear as the value in the primary key field, *sid,* of Students. Incidentally, the primary key constraint for Enrolled states that a student has exactly one grade for each course he or she is enrolled in. If we want to record more than one grade per student per course, we should change the primary key constraint.

3.2.3 General Constraints

Domain, primary key, and foreign key constraints are considered to be a fundamental part of the relational data model and are given special attention in most commercial systems. Sometimes, however, it is necessary to specify more general constraints.

For example, we may require that student ages be within a certain range of values; given such an IC specification, the DBMS rejects inserts and updates that violate the constraint. This is very useful in preventing data entry errors. If we specify that all students must be at least 16 years old, the instance of Students shown in Figure 3.1 is illegal because two students are underage. If we disallow the insertion of these two tuples, we have a legal instance, as shown in Figure 3.5.

sid	*name*	*login*	*age*	*gpa*
53666	Jones	jones@cs	18	3.4
53688	Smith	smith@ee	18	3.2
53650	Smith	smith@math	19	3.8

Figure 3.5 An Instance 82 of the Students Relation

The IC that students must be older than 16 can be thought of as an extended domain constraint, since we are essentially defining the set of permissible *age*

values more stringently than is possible by simply using a standard domain such as integer. In general, however, constraints that go well beyond domain, key, or foreign key constraints can be specified. For example, we could require that every student whose age is greater than 18 must have a gpa greater than 3.

Current relational database systems support such general constraints in the form of *table constraints* and *assertions*. Table constraints are associated with a single table and checked whenever that table is modified. In contrast, assertions involve several tables and are checked whenever any of these tables is modified. Both table constraints and assertions can use the full power of SQL queries to specify the desired restriction. We discuss SQL support for *table constraints* and *assertions* in Section 5.7 because a full appreciation of their power requires a good grasp of SQL's query capabilities.

3.3 ENFORCING INTEGRITY CONSTRAINTS

As we observed earlier, ICs are specified when a relation is created and enforced when a relation is modified. The impact of domain, PRIMARY KEY, and UNIQUE constraints is straightforward: If an insert, delete, or update command causes a violation, it is rejected. Every potential Ie violation is generally checked at the end of each SQL statement execution, although it can be *deferred* until the end of the transaction executing the statement, as we will see in Section 3.3.1.

Consider the instance 51 of Students shown in Figure 3.1. The following insertion violates the primary key constraint because there is already a tuple with the *s'id* 53688, and it will be rejected by the DBMS:

```
INSERT
INTO    Students   (sid, name, login, age, gpa)
VALUES  (53688, 'Mike', 'mike@ee', 17,3.4)
```

The following insertion violates the constraint that the primary key cannot contain *null*:

```
INSERT
INTO    Students   (sid, name, login, age, gpa)
VALUES  (null, 'Mike', 'mike@ee', 17,3.4)
```

Of course, a similar problem arises whenever we try to insert a tuple with a value in a field that is not in the domain associated with that field, that is, whenever we violate a domain constraint. Deletion does not cause a violation of clornain, primary key or unique constraints. However, an update can cause violations, sirnilar to an insertion:

```
UPDATE Students S
SET    S.sid = 50000
WHERE  S.sid = 53688
```

This update violates the primary key constraint because there is already a tuple with *sid 50000.*

The impact of foreign key constraints is more complex because SQL sometimes tries to rectify a foreign key constraint violation instead of simply rejecting the change. We discuss the referential integrity enforcement steps taken by the DBMS in terms of our Enrolled and Students tables, with the foreign key constraint that Enrolled.sid is a reference to (the primary key of) Students.

In addition to the instance 81 of Students, consider the instance of Enrolled shown in Figure 3.4. Deletions of Enrolled tuples do not violate referential integrity, but insertions of Enrolled tuples could. The following insertion is illegal because there is no Students tuple with *sid* 51111:

```
INSERT
INTO   Enrolled   (cid, grade, studid)
VALUES ('Hindi101', 'B', 51111)
```

On the other hand, insertions of Students tuples do not violate referential integrity, and deletions of Students tuples could cause violations. Further, updates on either Enrolled or Students that change the *studid* (respectively, *sid*) value could potentially violate referential integrity.

SQL provides several alternative ways to handle foreign key violations. We must consider three basic questions:

1. *What should we do if an Enrolled row is inserted, with a* studid *column value that does not appear in any row of the Students table?*

   In this case, the INSERT command is simply rejected.

2. *What should we do if a Students row is deleted?*

   The options are:

   - Delete all Enrolled rows that refer to the deleted Students row.
   - Disallow the deletion of the Students row if an Enrolled row refers to it.
   - Set the *studid* column to the *sid* of some (existing) 'default' student, for every Enrolled row that refers to the deleted Students row.

- For every Enrolled row that refers to it, set the *studid* column to *null*. In our example, this option conflicts with the fact that *stud'id* is part of the primary key of Enrolled and therefore cannot be set to *null*. Therefore, we are limited to the first three options in our example, although this fourth option (setting the foreign key to *null)* is available in general.

3. *What should we do if the primary key val'ue of a Students row* is *updated?* The options here are similar to the previous case.

SQL allows us to choose any of the four options on DELETE and UPDATE. For example, we can specify that when a Students row is *deleted,* all Enrolled rows that refer to it are to be deleted as well, but that when the *sid* column of a Students row is *modified,* this update is to be rejected if an Enrolled row refers to the modified Students row:

```
CREATE TABLE Enrolled ( studid CHAR(20),
                        cid    CHAR(20),
                        grade  CHAR(10),
                        PRIMARY KEY (studid, dd),
                        FOREIGN KEY (studid) REFERENCES Students
                                ON DELETE CASCADE
                                ON UPDATE NO ACTION)
```

The options are specified as part of the foreign key declaration. The default option is NO ACTION, which means that the action (DELETE or UPDATE) is to be rejected, Thus, the ON UPDATE clause in our example could be omitted, with the same effect. The CASCADE keyword says that, if a Students row is deleted, all Enrolled rows that refer to it are to be deleted as well. If the UPDATE clause specified CASCADE, and the *sid* column of a Students row is updated, this update is also carried out in each Enrolled row that refers to the updated Students row.

If a Students row is deleted, we can switch the enrollment to a 'default' student by using ON DELETE SET DEFAULT. The default student is specified as part of the definition of the *sid* field in Enrolled; for example, *sid* CHAR(20) DEFAULT *'53666'*. Although the specification of a default value is appropriate in some situations (e.g" a default parts supplier if a particular supplier goes out of business), it is really not appropriate to switch enrollments to a default student. The correct solution in this example is to also delete all enrollment tuples for the deleted student (that is, CASCADE) or to reject the update.

SQL also allows the use of *null* as the default value by specifying ON DELETE SET NULL.

3.3.1 Transactions and Constraints

As we saw in Chapter 1, a program that runs against a database is called a transaction, and it can contain several statements (queries, inserts, updates, etc.) that access the database. If (the execution of) a statement in a transaction violates an integrity constraint, should the DBMS detect this right away or should all constraints be checked together just before the transaction completes?

By default, a constraint is checked at the end of every SQL statement that could lead to a violation, and if there is a violation, the statement is rejected. Sometimes this approach is too inflexible. Consider the following variants of the Students and Courses relations; every student is required to have an honors course, and every course is required to have a grader, who is some student.

```
CREATE TABLE Students ( sid     CHAR(20),
                        name    CHAR(30),
                        login   CHAR(20),
                        age     INTEGER,
                        honors  CHAR(10) NOT NULL,
                        gpa     REAL)
                        PRIMARY KEY (sid),
                        FOREIGN KEY (honors) REFERENCES Courses (cid))

CREATE TABLE Courses (cid     CHAR(10),
                      cname   CHAR(10),
                      credits INTEGER,
                      grader  CHAR(20) NOT NULL,
                      PRIMARY KEY (dd)
                      FOREIGN KEY (grader) REFERENCES Students (sid))
```

vVhenever a Students tuple is inserted, a check is made to see if the"honors course is in the Courses relation, and whenever a Courses tuple is inserted, a check is made to see that the grader is in the Students relation. How are we to insert the very first course or student tuple? One cannot be inserted without the other. The only way to accomplish this insertion is to defer the constraint checking that would normally be carried out at the end of an INSERT statement.

SQL allows a constraint to be in DEFERRED or IMMEDIATE mode.

```
SET CONSTRAINT ConstraintFoo DEFERRED
```

A constraint in deferred mode is checked at commit time. In our example, the foreign key constraints on Boats and Sailors can both be declared to be in deferred mode. We can then insert a boat with a nonexistent sailor as the captain (temporarily making the database inconsistent), insert the sailor (restoring consistency), then commit and check that both constraints are satisfied.

3.4 QUERYING RELATIONAL DATA

A relational database query (query, for short) is a question about the data, and the answer consists of a new relation containing the result. For example, we might want to find all students younger than 18 or all students enrolled in Reggae203. A query language is a specialized language for writing queries.

SQL is the most popular commercial query language for a relational DBMS. We now present some SQL examples that illustrate how easily relations can be queried. Consider the instance of the Students relation shown in Figure 3.1. We can retrieve rows corresponding to students who are younger than 18 with the following SQL query:

```
SELECT *
FROM   Students S
WHERE  S.age < 18
```

The symbol ,*, means that we retain all fields of selected tuples in the result. Think of S as a variable that takes on the value of each tuple in Students, one tuple after the other. The condition *S.age* < 18 in the WHERE clause specifies that we want to select only tuples in which the *age* field has a value less than 18. This query evaluates to the relation shown in Figure 3.6.

sid	*name*	*login*	*age*	*gpa*
53831	Madayan	madayan@music	11	1.8
53832	Guldu	guldu@music	12	2.0

Figure 3.6 Students with *age* < 18 on Instance 51

This example illustrates that the domain of a field restricts the operations that are permitted on field values, in addition to restricting the values that can appear in the field. The condition *S.age* < 18 involves an arithmetic comparison of an *age* value with an integer and is permissible because the domain of *age* is the set of integers. On the other hand, a condition such as *S.age = S."id* does not make sense because it compares an integer value with a string value, and this comparison is defined to fail in SQL; a query containing this condition produces no answer tuples.

In addition to selecting a subset of tuples, a query can extract a subset of the fields of each selected tuple. We can compute the names and logins of students who are younger than 18 with the following query:

```
SELECT S.name, S.login
FROM   Students S
WHERE  S.age < 18
```

Figure 3.7 shows the answer to this query; it is obtained by applying the selection to the instance 81 of Students (to get the relation shown in Figure 3.6), followed by removing unwanted fields. Note that the order in which we perform these operations does matter-if we remove unwanted fields first, we cannot check the condition *S.age* < 18, which involves one of those fields.

name	*login*
Madayan	madayan@music
Guldu	guldu@music

Figure 3.7 Names and Logins of Students under 18

We can also combine information in the Students and Enrolled relations. If we want to obtain the names of all students who obtained an A and the id of the course in which they got an A, we could write the following query:

```
SELECT S.name, E.cid
FROM   Students S, Enrolled E
WHERE  S.sid = E.studid AND E.grade = 'A'
```

This query can be understood as follows: "If there is a Students tuple S and an Enrolled tuple E such that S.sid = E.studid (so that S describes the student who is enrolled in E) and E.grade = 'A', then print the student's name and the course id." When evaluated on the instances of Students and Enrolled in Figure 3.4, this query returns a single tuple, *(Smith, Topology112)*.

We cover relational queries and SQL in more detail in subsequent chapters.

3.5 LOGICAL DATABASE DESIGN: ER TO RELATIONAL

The ER model is convenient for representing an initial, high-level database design. Given an ER diagram describing a database, a standard approach is taken to generating a relational database schema that closely approximates

the ER design. (The translation is approximate to the extent that we cannot capture all the constraints implicit in the ER design using SQL, unless we use certain SQL constraints that are costly to check.) We now describe how to translate an ER diagram into a collection of tables with associated constraints, that is, a relational database schema.

3.5.1 Entity Sets to Tables

An entity set is mapped to a relation in a straightforward way: Each attribute of the entity set becomes an attribute of the table. Note that we know both the domain of each attribute and the (primary) key of an entity set.

Consider the Employees entity set with attributes *ssn, name,* and *lot* shown in Figure 3.8. A possible instance of the Employees entity set, containing three

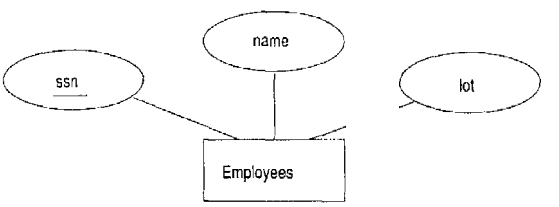


Figure 3.8 The Employees Entity Set

Employees entities, is shown in Figure 3.9 in a tabular format.

ssn	*name*	*lot*
123-22-3666	Attishoo	48
231-31-5368	Smiley	22
131-24-3650	Smethurst	35

Figure 3.9 An Instance of the Employees Entity Set

The following SQL statement captures the preceding information, including the domain constraints and key information:

```
CREATE TABLE Employees ( ssn        CHAR(11),
                         name       CHAR(30) ,
                         lot        INTEGER,
                         PRIMARY KEY (ssn) )
```

3.5.2 Relationship Sets (without Constraints) to Tables

A relationship set, like an entity set, is mapped to a relation in the relational model. We begin by considering relationship sets without key and participation constraints, and we discuss how to handle such constraints in subsequent sections. To represent a relationship, we must be able to identify each participating entity and give values to the descriptive attributes of the relationship. Thus, the attributes of the relation include:

- The primary key attributes of each participating entity set, as foreign key fields.
- The descriptive attributes of the relationship set.

The set of nondescriptive attributes is a superkey for the relation. If there are no key constraints (see Section 2.4.1), this set of attributes is a candidate key.

Consider the Works_In2 relationship set shown in Figure 3.10. Each department has offices in several locations and we want to record the locations at which each employee works.

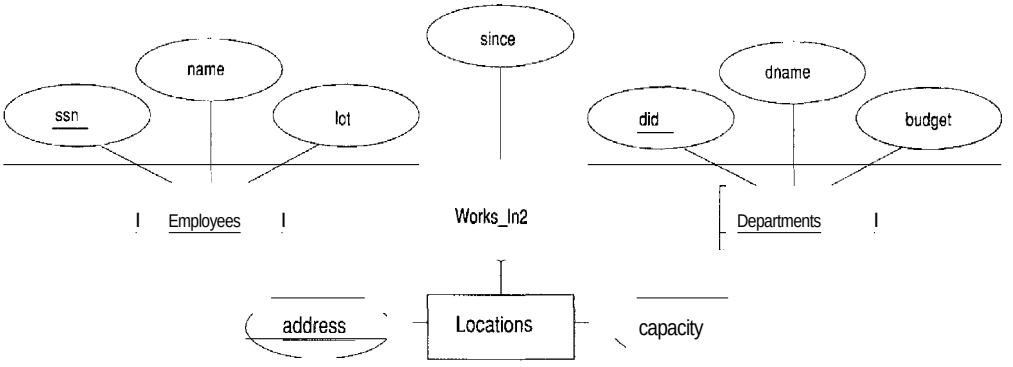


Figure 3.10 A Ternary Relationship Set

All the available information about the Works_In2 table is captured by the following SQL definition:

```
CREATE TABLE Works_In2 ( ssn       CHAR(11),
                         did       INTEGER,
                         address   CHAR(20),
                         since     DATE,
                         PRIMARY KEY (8sn, did, address),
                         FOREIGN KEY (ssn) REFERENCES Employees,
```

```
                    FOREIGN KEY (address) REFERENCES Locations,
                    FOREIGN KEY (did) REFERENCES Departments)
```

Note that the *address, did.* and *ssn* fields cannot take on *n'ull* values. Because these fields are part of the primary key for \Vorks_In2, a NOT NULL constraint is implicit for each of these fields. This constraint ensures that these fields uniquely identify a department, an employee, and a location in each tuple of WorksJn. We can also specify that a particular action is desired when a referenced Employees, Departments, or Locations tuple is deleted, as explained in the discussion of integrity constraints in Section 3.2. In this chapter, we assume that the default action is appropriate except for situations in which the semantics of the ER diagram require some other action.

Finally, consider the Reports_To relationship set shown in Figure 3.11. The

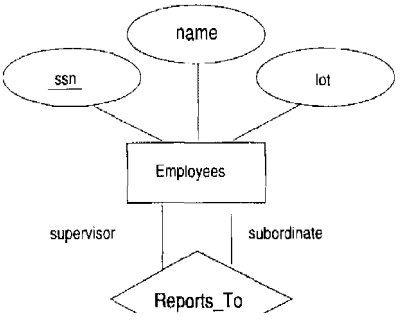


Figure 3.11 The Reports_To Relationship Set

role indicators *supervisor* and *subordinate* are used to create meaningful field names in the CREATE statement for the Reports_To table:

```
CREATE TABLE Reports_To (
        supervisor...ssn    CHAR(11),
        subordinate...ssn   CHAR(11),
        PRIMARY KEY (supervisor_ssn, subordinate_ssn),
        FOREIGN KEY (supervisor...ssn) REFERENCES Employees(ssn),
        FOREIGN KEY (subordinate...ssn) REFERENCES Employees(ssn) )
```

Observe that we need to explicitly name the referenced field of Employees because the field name differs from the name(s) of the referring field(s).

3.5.3 Translating Relationship Sets with Key Constraints

If a relationship set involves n entity sets and somem of them are linked via arrows in the ER diagTam, the key for anyone of these m entity sets constitutes a key for the relation to which the relationship set is mapped. Hence we have m candidate keys, and one of these should be designated as the primary key. The translation discussed in Section 2.3 from relationship sets to a relation can be used in the presence of key constraints, taking into account this point about keys.

Consider the relationship set Manages shown in Figure 3.12. The table cor-

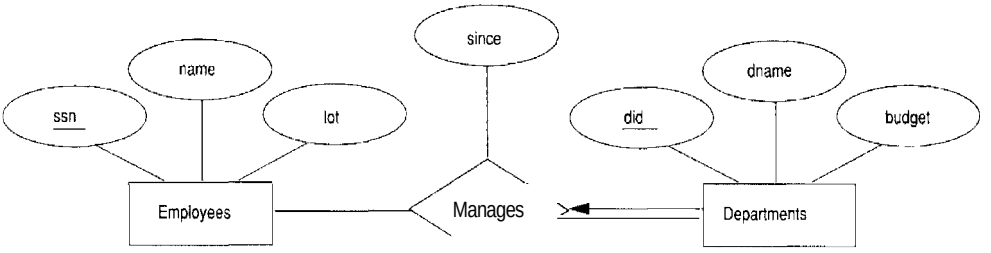


Figure 3.12 Key Constraint on Manages

responding to Manages has the attributes *ssn, did, since.* However, because each department has at most one manager, no two tuples can have the same *did* value but differ on the *ssn* value. A consequence of this observation is that *did* is itself a key for Manages; indeed, the set *did, ssn* is not a key (because it is not minimal). The Manages relation can be defined using the following SQL statement:

```
CREATE TABLE Manages (ssn      CHAR(11),
                      did      INTEGER,
                      since    DATE,
                      PRIMARY KEY (did),
                      FOREIGN KEY (ssn) REFERENCES Employees,
                      FOREIGN KEY (did) REFERENCES Departments)
```

A second approach to translating a relationship set with key constraints is often superior because it avoids creating a distinct table for the relationship set. The idea is to include the information about the relationship set in the table corresponding to the entity set with the key, taking adyantage of the key constraint. In the Manages example, because a department has at most one manager, we can add the key fields of the Employees tuple denoting the Inanager and the *since* attribute to the Departments tuple.

This approach eliminates the need for a separate Manages relation, and queries asking for a department's manager can be answered without combining information from two relations. The only drawback to this approach is that space could be wasted if several departments have no managers. In this case the added fields would have to be filled with *null* values. The first translation (using a separate table for Manages) avoids this inefficiency, but some important queries require us to combine information from two relations, which can be a slow operation.

The following SQL statement, defining a DepLMgr relation that captures the information in both Departments and Manages, illustrates the second approach to translating relationship sets with key constraints:

```
CREATE TABLE DepLMgr ( did        INTEGER,
                       dname      CHAR(20),
                       budget     REAL,
                       ssn        CHAR(11),
                       since      DATE,
                       PRIMARY KEY (did),
                       FOREIGN KEY (ssn) REFERENCES Employees)
```

Note that *ssn* can take on *null* values.

This idea can be extended to deal with relationship sets involving more than two entity sets. In general, if a relationship set involves *n* entity sets and some *m* of them are linked via arrows in the ER diagram, the relation corresponding to anyone of the m sets can be augmented to capture the relationship.

We discuss the relative merits of the two translation approaches further after considering how to translate relationship sets with participation constraints into tables.

3.5.4 Translating Relationship Sets with **Participation** Constraints

Consider the ER diagram in Figure 3.13, which shows two relationship sets, Manages and Works_In.

Every department is required to have a manager, due to the participation constraint, and at most one manager, due to the key constraint. The following SQL statement reflects the second translation approach discussed in Section 3.5.3, and uses the key constraint:

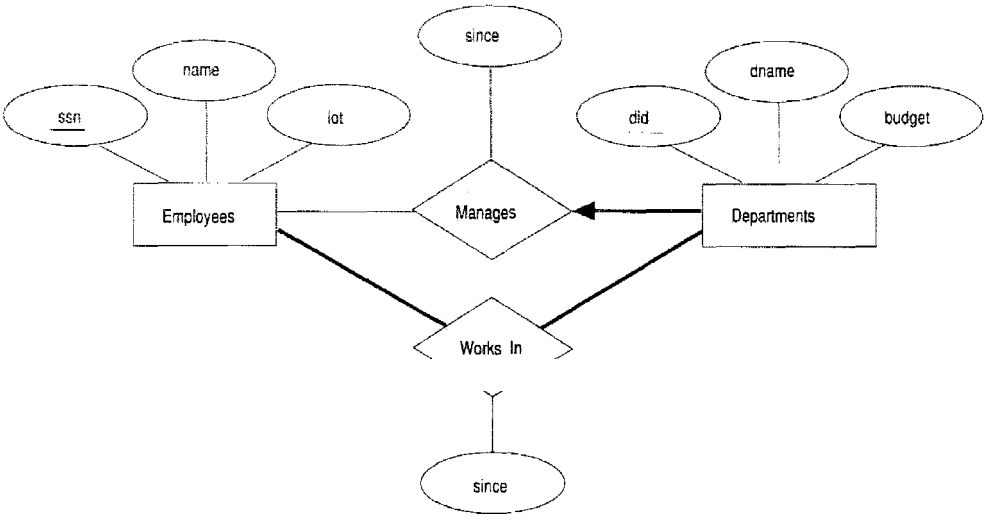


Figure 3.13 Manages and WorksJn

```
CREATE TABLE Dept_Mgr ( did       INTEGER,
                        dname     CHAR(20) ,
                        budget    REAL,
                        ssn       CHAR(11) NOT NULL,
                        since     DATE,
                        PRIMARY KEY (did),
                        FOREIGN KEY (ssn) REFERENCES Employees
                                ON DELETE NO ACTION)
```

It also captures the participation constraint that every department must have a manager: Because *ssn* cannot take on *null* values, each tuple of Dept_Mgr identifies a tuple in Employees (who is the manager). The NO ACTION specification, which is the default and need not be explicitly specified, ensures that an Employees tuple cannot be deleted while it is pointed to by a Dept_Mgr tuple. If we wish to delete such an Employees tuple, we must first change the Dept_Mgr tuple to have a new employee as manager. (We could have specified CASCADE instead of NO ACTION, but deleting all information about a department just because its manager has been fired seems a bit extreme!)

The constraint that every department must have a manager cannot be captured using the first translation approach discussed in Section 3.5.3. (Look at the definition of Manages and think about what effect it would have if we added NOT NULL constraints to the *ssn* and *did* fields. *Hint:* The constraint would prevent the firing of a manager, but does not ensure that a manager is initially appointed for each department!) This situation is a strong argument

in favor of using the second approach for one-to-many relationships such as Manages, especially when the entity set with the key constraint also has a total participation constraint.

Unfortunately, there are many participation constraints that we cannot capture using SQL, short of using *table constraints* or *assertions*. Table constraints and assertions can be specified using the full power of the SQL query language (as discussed in Section 5.7) and are very expressive but also very expensive to check and enforce. For example, we cannot enforce the participation constraints on the Works_In relation without using these general constraints. To see why, consider the Works_In relation obtained by translating the ER diagram into relations. It contains fields *ssn* and *did*, which are foreign keys referring to Employees and Departments. To ensure total participation of Departments in Works_In, we have to guarantee that every *did* value in Departments appears in a tuple of Works_In. We could try to guarantee this condition by declaring that *did* in Departments is a foreign key referring to Works_In, but this is not a valid foreign key constraint because *did* is not a candidate key for Works_In.

To ensure total participation of Departments in Works_In using SQL, we need an assertion. We have to guarantee that every *did* value in Departments appears in a tuple of Works_In; further, this tuple of Works_In must also have *non-null* values in the fields that are foreign keys referencing other entity sets involved in the relationship (in this example, the *ssn* field). We can ensure the second part of this constraint by imposing the stronger requirement that *ssn* in Works_In cannot contain *null* values. (Ensuring that the participation of Employees in Works_In is total is symmetric.)

Another constraint that requires assertions to express in SQL is the requirement that each Employees entity (in the context of the Manages relationship set) must manage at least one department.

In fact, the Manages relationship set exemplifies most of the participation constraints that we can capture using key and foreign key constraints. Manages is a binary relationship set in which exactly one of the entity sets (Departments) has a key constraint, and the total participation constraint is expressed on that entity set.

We can also capture participation constraints using key and foreign key constraints in one other special situation: a relationship set in which all participating entity sets have key constraints and total participation. The best translation approach in this case is to map all the entities as well as the relationship into a single table; the details are straightforward.

3.5.5 Translating Weak Entity Sets

A weak entity set always participates in a one-to-many binary relationship and has a key constraint and total participation. The second translation approach discussed in Section 3.5.3 is ideal in this case, but we must take into account that the weak entity has only a partial key. Also, when an owner entity is deleted, we want all owned weak entities to be deleted.

Consider the Dependents weak entity set shown in Figure 3.14, with partial key *pname*. A Dependents entity can be identified uniquely only if we take the key of the *owning* Employees entity and the *pname* of the Dependents entity, and the Dependents entity must be deleted if the owning Employees entity is deleted.

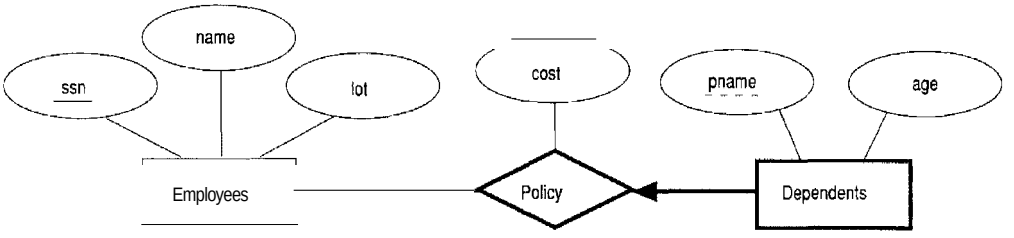


Figure 3.14 The Dependents Weak Entity Set

We can capture the desired semantics with the following definition of the Dep_Policy relation:

```
CREATE TABLE Dep_Policy (pname   CHAR(20),
                         age     INTEGER,
                         cost    REAL,
                         ssn     CHAR(11),
                         PRIMARY KEY (pname, ssn),
                         FOREIGN KEY (ssn) REFERENCES Employees
                                 ON DELETE CASCADE )
```

Observe that the primary key is *(pna:me, ssn)*, since Dependents is a weak entity. This constraint is a change with respect to the translation discussed in Section 3.5.3. \Ve have to ensure that every Dependents entity is associated with an Employees entity (the owner), as per the total participation constraint on Dependents. That is, *ssn* cannot be *null*. This is ensured because *SST,* is part of the primary key. The CASCADE option ensures that information about an employee's policy and dependents is deleted if the corresponding Employees tuple is deleted.

3.5.6 Translating Class Hierarchies

We present the two basic approaches to handling ISA hierarchies by applying them to the ER diagram shown in Figure 3.15:

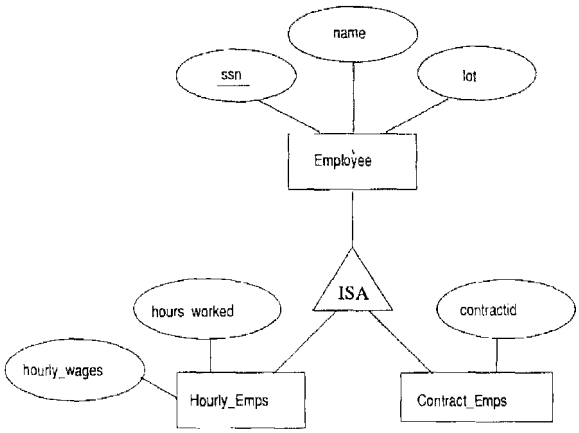


Figure 3.15 Class Hierarchy

1. We can map each of the entity sets Employees, Hourly_Emps, and ContracLEmps to a distinct relation. The Employees relation is created as in Section 2.2. We discuss Hourly_Emps here; ContracLEmps is handled similarly. The relation for Hourly_Emps includes the *hourly_wages* and *hours_worked* attributes of Hourly_Emps. It also contains the key attributes of the superclass *(ssn,* in this example), which serve as the primary key for Hourly_Emps, as well as a foreign key referencing the superclass (Employees). For each Hourly_Emps entity, the value of the *name* and *lot* attributes are stored in the corresponding row of the superclass (Employees). Note that if the superclass tuple is deleted, the delete must be cascaded to Hourly_Emps.

2. Alternatively, we can create just two relations, corresponding to Hourly_Emps and ContracLEmps. The relation for Hourly_Emps includes all the attributes of Hourly_Emps as well as all the attributes of Employees (i.e., *ssn, name, lot, hourly_wages, hours_worked*).

The first approach is general and always applicable. Queries in which we want to examine all employees and do not care about the attributes specific to the subclasses are handled easily using the Employees relation. However, queries in which we want to examine, say, hourly employees, may require us to combine Hourly_Emps (or ContracLEmps, as the case may be) with Employees to retrieve *name* and *lot.*

The second approach is not applicable if we have employees who are neither hourly employees nor contract employees, since there is no way to store such employees. Also, if an employee is both an Hourly_Emps and a ContracLEmps entity, then the *name* and *lot* values are stored twice. This duplication can lead to some of the anomalies that we discuss in Chapter 19. A query that needs to examine all employees must now examine two relations. On the other hand, a query that needs to examine only hourly employees can now do so by examining just one relation. The choice between these approaches clearly depends on the semantics of the data and the frequency of common operations.

In general, overlap and covering constraints can be expressed in SQL only by using assertions.

3.5.7 Translating ER Diagrams with Aggregation

Consider the ER diagram shown in Figure 3.16. The Employees, Projects,

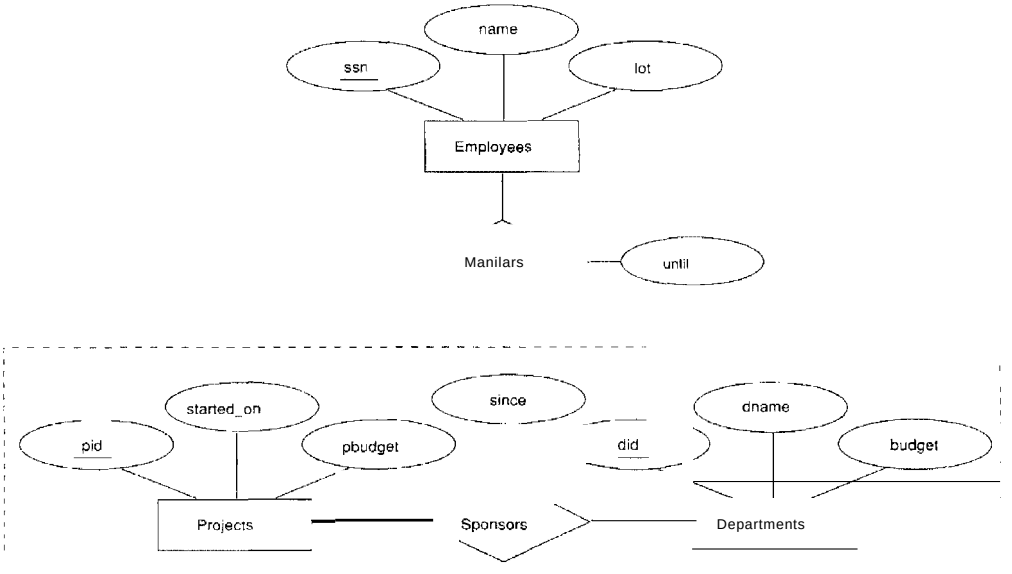


Figure 3.16 Aggregation

and Departments entity sets and the Sponsors relationship set are mapped as described in previous sections. For the Monitors relationship set, we create a relation with the following attributes: the key attributes of Employees *(88n),* the key attributes of Sponsors (did, pid), and the descriptive attributes of Monitors ($until$). This translation is essentially the standard mapping for a relationship set, as described in Section 3.5.2.

There is a special case in which this translation can be refined by dropping the Sponsors relation. Consicler the Sponsors relation. It has attributes *pid, did,* and *since;* and in general we need it (in addition to l\rlonitors) for two reasons:

1. \Ve have to record the descriptive attributes (in our example, *since)* of the Sponsors relationship.
2. Not every sponsorship has a monitor, and thus some $\langle pid, did \rangle$ pairs in the Sponsors relation may not appear in the Monitors relation.

However, if Sponsors has no descriptive attributes and has total participation in Monitors, every possible instance of the Sponsors relation can be obtained from the $\langle pid, did \rangle$ columns of Monitors; Sponsors can be dropped.

3.5.8 ER to Relational: Additional Examples

Consider the ER diagram shown in Figure 3.17. We can use the key constraints

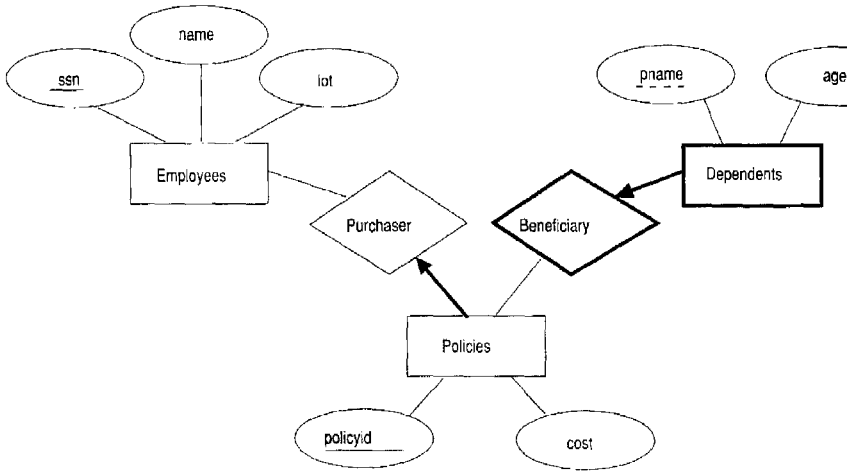


Figure 3.17 Policy Revisited

to combine Purchaser information with Policies and Beneficiary information with Dependents, and translate it into the relational model as follows:

```
CREATE TABLE Policies ( policyid INTEGER,
                        cost     REAL,
                        ssn      CHAR(11) NOT NULL,
                        PRIMARY KEY (policyid),
                        FOREIGN KEY (ssn) REFERENCES Employees
                                ON DELETE CASCADE )
```

```
CREATE TABLE Dependents (pname    CHAR(20),
                         age      INTEGER,
                         policyid INTEGER,
                         PRIMARY KEY (pname, policyid),
                         FOREIGN KEY (policyid) REFERENCES Policies
                                 ON DELETE CASCADE)
```

Notice how the deletion of an employee leads to the deletion of all policies owned by the employee and all dependents who are beneficiaries of those policies. Further, each dependent is required to have a covering policy-because *policyid* is part of the primary key of Dependents, there is an implicit NOT NULL constraint. This model accurately reflects the participation constraints in the ER diagram and the intended actions when an employee entity is deleted.

In general, there could be a chain of identifying relationships for weak entity sets. For example, we assumed that *policyid* uniquely identifies a policy. Suppose that *policyid* distinguishes only the policies owned by a given employee; that is, *policyid* is only a partial key and Policies should be modeled as a weak entity set. This new assumption about *policyid* does not cause much to change in the preceding discussion. In fact, the only changes are that the primary key of Policies becomes *(policyid, ssn)*, and as a consequence, the definition of Dependents changes-a field called *ssn* is added and becomes part of both the primary key of Dependents and the foreign key referencing Policies:

```
CREATE TABLE Dependents (pname    CHAR(20),
                         ssn      CHAR(11),
                         age      INTEGER,
                         policyid INTEGER NOT NULL,
                         PRIMARY KEY (pname, policyid, ssn),
                         FOREIGN KEY (policyid, ssn) REFERENCES Policies
                                 ON DELETE CASCADE )
```

3.6 INTRODUCTION TO VIEWS

A view is a table whose rows are not explicitly stored in the database but are computed as needed from a view definition. Consider the Students and Enrolled relations. Suppose we are often interested in finding the names and student identifiers of students who got a grade of B in some course, together with the course identifier. We can define a view for this purpose. Using SQL notation:

```
CREATE VIEW B-Students (name, sid, course)
        AS SELECT S.sname, S.sid, E.cid
```

```
FROM    Students S, Enrolled E
WHERE   S.sid = E.studid AND E.grade = 'B'
```

The view B-Students has three fields called *name, sid,* and *course* with the same domains as the fields *sname* and *sid* in *Students* and *cid* in Enrolled. (If the optional arguments *name, sid,* and *course* are omitted from the CREATE VIEW statement, the column names *sname, sid,* and *cid* are inherited.)

This view can be used just like a base table, or explicitly stored table, in defining new queries or views. Given the instances of Enrolled and Students shown in Figure 3.4, B-Students contains the tuples shown in Figure 3.18. Conceptually, whenever B-Students is used in a query, the view definition is first evaluated to obtain the corresponding instance of B-Students, then the rest of the query is evaluated treating B-Students like any other relation referred to in the query. (We discuss how queries on views are evaluated in practice in Chapter 25.)

name	*sid*	*course*
Jones	53666	History105
Guldu	53832	Reggae203

Figure 3.18 An Instance of the B-Students View

3.6.1 Views, Data Independence, Security

Consider the levels of abstraction we discussed in Section 1.5.2. The *physical* schema for a relational database describes how the relations in the conceptual schema are stored, in terms of the file organizations and indexes used. The *conceptual* schema is the collection of schemas of the relations stored in the database. While some relations in the conceptual schema can also be exposed to applications, that is, be part of the *exte'mal* schema of the database, additional relations in the *external* schema can be defined using the view mechanism. The view mechanism thus provides the support for *logical data independence* in the relational model. That is, it can be used to define relations in the external schema that mask changes in the conceptual schema of the database from applications. For example, if the schema of a stored relation is changed, we can define a view with the old schema and applications that expect to see the old schema can now use this view.

Views are also valuable in the context of *security:* We can define views that give a group of users access to just the information they are allowed to see. For example, we can define a view that allows students to see the other students'

name and age but not their gpa, and allows all students to access this view but not the underlying Students table (see Chapter 21).

3.6.2 Updates on Views

The motivation behind the view mechanism is to tailor how users see the data. Users should not have to worry about the view versus base table distinction. This goal is indeed achieved in the case of queries on views; a view can be used just like any other relation in defining a query. However, it is natural to want to specify updates on views as well. Here, unfortunately, the distinction between a view and a base table must be kept in mind.

The SQL-92 standard allows updates to be specified only on views that are defined on a single base table using just selection and projection, with no use of aggregate operations.[3] Such views are called updatable views. This definition is oversimplified, but it captures the spirit of the restrictions. An update on such a restricted view can always be implemented by updating the underlying base table in an unambiguous way. Consider the following view:

```
CREATE VIEW GoodStudents (sid, gpa)
        AS SELECT S.sid, S.gpa
           FROM   Students S
           WHERE  S.gpa> 3.0
```

We can implement a command to modify the gpa of a GoodStudents row by modifying the corresponding row in Students. We can delete a GoodStudents row by deleting the corresponding row from Students. (In general, if the view did not include a key for the underlying table, several rows in the table could 'correspond' to a single row in the view. This would be the case, for example, if we used *S.sname* instead of *S.sid* in the definition of GoodStudents. A command that affects a row in the view then affects all corresponding rows in the underlying table.)

We can insert a GoodStudents row by inserting a row into Students, using *null* values in columns of Students that do not appear in GoodStudents (e.g., *sname, login).* Note that primary key columns are not allowed to contain *null* values. Therefore, if we attempt to insert rows through a view that does not contain the primary key of the underlying table, the insertions will be rejected. For example, if GoodStudents contained *sname* but not *sid*, we could not insert rows into Students through insertions to GooclStudents.

3There is also the restriction that the DISTINCT operator cannot be used in updatable view definitions. By default, SQL does not eliminate duplicate copies of rows from the result of a query; the DISTINCT operator requires duplicate elimination. We discuss this point further in Chapter 5.

Updatable Views in SQL:1999 The Hew SQL standard has expanded the class of view definitions that are updatable, taking primary key constraints into account. In contrast to SQL-92, a view definition that contains more than Olle table in the FROM clause may be updatable under the new definition. Intuitively, we can update afield of a view if it is obtained from exactly one of the underlying tables, and the primary key of that table is included in the fields of the view.

SQL:1999 distinguishes between views whose rows can be modified *(updatable views)* and views into which new rows can be inserted (insertable-into views): Views defined using the SQL constructs UNION, INTERSECT, and EXCEPT (which we discuss in Chapter 5) cannot be inserted into, even if they are updatable. Intuitively, updatability ensures that an updated tuple in the view can be traced to exactly one tuple in one of the tables used to define the view. The updatability property, however, may still not enable us to decide into which table to insert a new tuple.

An important observation is that an INSERT or UPDATE may change the underlying base table so that the resulting (i.e., inserted or modified) row is not in the view! For example, if we try to insert a row (51234, 2.8) into the view, this row can be (padded with *null* values in the other fields of Students and then) added to the underlying Students table, but it will not appear in the GoodStudents view because it does not satisfy the view condition *gpa* > 3.0. The SQL default action is to allow this insertion, but we can disallow it by adding the clause WITH CHECK OPTION to the definition of the view. In this case, only rows that will actually appear in the view are permissible insertions.

We caution the reader, that when a view is defined in terms of another view, the interaction between these view definitions with respect to updates and the CHECK OPTION clause can be complex; we not go into the details.

Need to Restrict View Updates

vVhile the SQL rules on updatable views are more stringent than necessary, there are some fundamental problems with updates specified on views and good reason to limit the class of views that can be updated. Consider the Students relation and a new relation called Clubs:

Clubs(*cname:* string, *jyear:* date, *mname:* string)

	jyear	
Sailing	1996	Dave
Hiking	1997	Smith
Rowing	1998	Smith

Figure 3.19 An Instance C of Clubs

sid	*name*	*login*	*age*	*gpa*
50000	Dave	dave@cs	19	3.3
53666	Jones	jones@cs	18	3.4
53688	Smith	smith@ee	18	3.2
53650	Smith	smith@math	19	3.8

Figure 3.20 An Instance 53 of Students

name	*login*	*club*	*since*
Dave	dave@cs	Sailing	1996
Smith	smith@ee	Hiking	1997
Smith	smith@ee	Rowing	1998
Smith	smith@math	Hiking	1997
Smith	smith@math	Rowing	1998

Figure 3.21 Instance of ActiveStudents

A tuple in Clubs denotes that the student called *mname* has been a member of the club *cname* since the date *jyear.*[4] Suppose that we are often interested in finding the names and logins of students with a gpa greater than 3 who belong to at least one club, along with the club name and the date they joined the club. We can define a view for this purpose:

```
CREATE VIEW ActiveStudents (name, login, club, since)
        AS SELECT S.sname, S.login, C.cname, C.jyear
           FROM   Students S, Clubs C
           WHERE  S.sname = C.mname AND S.gpa> 3
```

Consider the instances of Students and Clubs shown in Figures 3.19 and 3.20. When evaluated using the instances C and *S3,* ActiveStudents contains the rows shown in Figure 3.21.

Now suppose that we want to delete the row *(Smith, smith@ee, Hiking, 1997)* from ActiveStudents. How are we to do this? ActiveStudents rows are not stored explicitly but computed as needed from the Students and Clubs tables using the view definition. So we must change either Students or Clubs (or both) in such a way that evaluating the view definition on the modified instance does not produce the row ⟨*Smith, smith@ee, Hiking,* 1997.) This task can be accomplished in one of two ways: by either deleting the row ⟨*53688, Smith, smith@ee,* 18, *3.2*⟩ from Students or deleting the row *(Hiking, 1997, Smith)*

[4]We remark that Clubs has a poorly designed schema (chosen for the sake of our discussion of view updates), since it identifies students by name, which is not a candidate key for Students.

from Clubs. But neither solution is satisfactory. Removing the Students row has the effect of also deleting the row *(8m:ith, smith@ee, Rowing,* 1998) from the view ActiveStudents. Removing the Clubs row has the effect of also deleting the row *(Smith, smith@math, Hiking, 1997)* from the view ActiveStudents. Neither side effect is desirable. In fact, the only reasonable solution is to *d'isallow* such updates on views.

Views involving more than one base table can, in principle, be safely updated. The B-Students view we introduced at the beginning of this section is an example of such a view. Consider the instance of B-Students shown in Figure 3.18 (with, of course, the corresponding instances of Students and Enrolled as in Figure 3.4). To insert a tuple, say *(Dave, 50000, Reggae203)* B-Students, we can simply insert a tuple *(Reggae203, B, 50000)* into Enrolled since there is already a tuple for *sid* 50000 in Students. To insert *(John, 55000, Reggae203),* on the other hand, we have to insert *(Reggae203, B, 55000)* into Enrolled and also insert *(55000, John,* null, null, null) into Students. Observe how *null* values are used in fields of the inserted tuple whose value is not available. Fortunately, the view schema contains the primary key fields of both underlying base tables; otherwise, we would not be able to support insertions into this view. To delete a tuple from the view B-Students, we can simply delete the corresponding tuple from Enrolled.

Although this example illustrates that the SQL rules on updatable views are unnecessarily restrictive, it also brings out the complexity of handling view updates in the general case. For practical reasons, the SQL standard has chosen to allow only updates on a very restricted class of views.

3.7 DESTROYING/ALTERING TABLES AND VIEWS

If we decide that we no longer need a base table and want to destroy it (i.e., delete all the rows *and* remove the table definition information), we can use the DROP TABLE command. For example, DROP TABLE Students RESTRICT destroys the Students table unless some view or integrity constraint refers to Students; if so, the command fails. If the keyword RESTRICT is replaced by CASCADE, Students is dropped and any referencing views or integrity constraints are (recursively) dropped as well; one of these two keywords must always be specified. A view can be dropped using the DROP VIEW command, which is just like DROP TABLE.

ALTER TABLE modifies the structure of an existing table. To add a column called *maiden-name* to Students, for example, we would use the following command:

```
ALTER TABLE Students
        ADD COLUMN maiden-name CHAR(10)
```

The definition of Students is modified to add this column, and all existing rows are padded with *null* values in this column. ALTER TABLE can also be used to delete columns and add or drop integrity constraints on a table; we do not discuss these aspects of the command beyond remarking that dropping columns is treated very similarly to dropping tables or views.

3.8 CASE STUDY: THE INTERNET STORE

The next design step in our running example, continued from Section 2.8, is logical database design. Using the standard approach discussed in Chapter 3, DBDudes maps the ER diagram shown in Figure 2.20 to the relational model, generating the following tables:

```
CREATE TABLE Books ( isbn            CHAR(10),
                     title           CHAR(80),
                     author          CHAR(80),
                     qty_in_stock    INTEGER,
                     price           REAL,
                     yeaLpublished   INTEGER,
                     PRIMARY KEY (isbn))

CREATE TABLE Orders ( isbn           CHAR(10),
                      ciel           INTEGER,
                      carelnum       CHAR(16),
                      qty            INTEGER,
                      order_date     DATE,
                      ship_date      DATE,
                      PRIMARY KEY (isbn,cid),
                      FOREIGN KEY (isbn) REFERENCES Books,
                      FOREIGN KEY (cid) REFERENCES Customers)

CREATE TABLE Customers ( cid         INTEGER,
                         cname       CHAR(80),
                         address     CHAR(200),
                         PRIMARY KEY (cid)
```

The design team leader, who is still brooding over the fact that the review exposed a flaw in the design, now has an inspiration. The Orders table contains the field *order_date* and the key for the table contains only the fields *isbn* and *c'id.* Because of this, a customer cannot order the same book on different days,

a restriction that was not intended. Why not add the *order_date* attribute to the key for the Orders table? This would eliminate the unwanted restriction:

```
CREATE TABLE Orders (   isbn          CHAR(10),

                        PRIMARY KEY (isbn,cid,ship_date),
                        ...)
```

The reviewer, Dude 2, is not entirely happy with this solution, which he calls a 'hack'. He points out that no natural ER diagram reflects this design and stresses the importance of the ER diagram as a design document. Dude 1 argues that, while Dude 2 has a point, it is important to present B&N with a preliminary design and get feedback; everyone agrees with this, and they go back to B&N.

The owner of B&N now brings up some additional requirements he did not mention during the initial discussions: "Customers should be able to purchase several different books in a single order. For example, if a customer wants to purchase three copies of 'The English Teacher' and two copies of 'The Character of Physical Law,' the customer should be able to place a single order for both books."

The design team leader, Dude 1, asks how this affects the shippping policy. Does B&N still want to ship all books in an order together? The owner of B&N explains their shipping policy: "As soon as we have have enough copies of an ordered book we ship it, even if an order contains several books. So it could happen that the three copies of 'The English Teacher' are shipped today because we have five copies in stock, but that 'The Character of Physical Law' is shipped tomorrow, because we currently have only one copy in stock and another copy arrives tomorrow. In addition, my customers could place more than one order per day, and they want to be able to identify the orders they placed."

The DBDudes team thinks this over and identifies two new requirements: First, it must be possible to order several different books in a single order and second, a customer must be able to distinguish between several orders placed the same day. To accomodate these requirements, they introduce a new attribute into the Orders table called *ordernum*, which uniquely identifies an order and therefore the customer placing the order. However, since several books could be purchased in a single order, *ordernum* and *isbn* are both needed to determine *qty* and *ship_date* in the Orders table.

Orders are assigned order numbers sequentially and orders that are placed later have higher order numbers. If several orders are placed by the same customer

on a single day, these orders have different order numbers and can thus be distinguished. The SQL DDL statement to create the modified Orders table follows:

```
CREATE TABLE Orders ( ordernum     INTEGER,
                      isbn         CHAR(10),
                      dd           INTEGER,
                      cardnum      CHAR(16),
                      qty          INTEGER,
                      ordeLdate    DATE,
                      ship_date    DATE,
                      PRIMARY KEY (ordernum, isbn),
                      FOREIGN KEY (isbn) REFERENCES Books
                      FOREIGN KEY (dd) REFERENCES Customers)
```

The owner of B&N is quite happy with this design for Orders, but has realized something else. (DBDudes is not surprised; customers almost always come up with several new requirements as the design progresses.) While he wants all his employees to be able to look at the details of an order, so that they can respond to customer enquiries, he wants customers' credit card information to be secure. To address this concern, DBDudes creates the following view:

```
CREATE VIEW OrderInfo (isbn, cid, qty, order-date, ship_date)
       AS SELECT O.cid, O.qty, O.ordeLdate, O.ship_date
          FROM   Orders 0
```

The plan is to allow employees to see this table, but not Orders; the latter is restricted to B&N's Accounting division. We'll see how this is accomplished in Section 21.7.

3.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What is a relation? Differentiate between a relation schema and a relation instance. Define the terms *arity* and *degree* of a relation. What are domain constraints? (Section 3.1)
- What SQL construct enables the definition of a relation? What constructs allow modification of relation instances? (Section 3.1.1)
- What are *integrity constraints?* Define the terms *primary key constraint* and *foreign key constraint.* How are these constraints expressed in SQL? What other kinds of constraints can we express in SQL? (Section 3.2)

- \Vhat does the DBMS do when constraints are violated? What is *referential 'integr-ity?* \Vhat options does SQL give application programmers for dealing with violations of referential integrity? (Section 3.3)
- When are integrity constraints enforced by a DBMS? How can an application programmer control the time that constraint violations are checked during transaction execution? (Section 3.3.1)
- What is a *relational database query?* (Section 3.4)
- How can we translate an ER diagram into SQL statements to create tables? How are entity sets mapped into relations? How are relationship sets mapped? How are constraints in the ER model, weak entity sets, class hierarchies, and aggregation handled? (Section 3.5)
- What is a *view?* How do views support logical data independence? How are views used for security? How are queries on views evaluated? Why does SQL restrict the class of views that can be updated? (Section 3.6)
- What are the SQL constructs to modify the structure of tables and destray tables and views? Discuss what happens when we destroy a view. (Section 3.7)

EXERCISES

Exercise 3.1 Define the following terms: *relation schema, relational database schema, domain, relation instance, relation cardinality,* and *relation degree.*

Exercise 3.2 How many distinct tuples are in a relation instance with cardinality 22?

Exercise 3.3 Does the relational model, as seen by an SQL query writer, provide physical and logical data independence? Explain.

Exercise 3.4 What is the difference between a candidate key and the primary key for a given relation? What is a superkey?

Exercise 3.5 Consider the instance of the Students relation shown in Figure 3.1.

1. Give an example of an attribute (or set of attributes) that you can deduce is *not* a candidate key, based on this instance being legaL
2. Is there any example of an attribute (or set of attributes) that you can deduce is a candidate key, based on this instance being legal?

Exercise 3.6 What is a foreign key constraint? Why are such constraints important? What is referential integrity?

Exercise 3.7 Consider the relations Students, Faculty, Courses, Rooms, Enrolled, Teaches, and Meets_In defined in Section 1.5.2.

1. List all the foreign key constraints among these relations.
2. Give an example of a (plausible) constraint involving one or more of these relations that is not a primary key or foreign key constraint.

Exercise 3.8 Answer each of the following questions briefly. The questions are based on the following relational schema:

Emp(*eid:* integer, *ename:* string, *age:* integer, *salary:* real)
Works(*eid:* integer, *did:* integer, *pct_time:* integer)
Dept(did: integer, *dname:* string, *budget:* real, *managerid:* integer)

1. Give an example of a foreign key constraint that involves the Dept relation. What are the options for enforcing this constraint when a user attempts to delete a Dept tuple?
2. Write the SQL statements required to create the preceding relations, including appropriate versions of all primary and foreign key integrity constraints.
3. Define the Dept relation in SQL so that every department is guaranteed to have a manager.
4. Write an SQL statement to add John Doe as an employee with *eid* = 101, *age* = 32 and *salary* = 15,000.
5. Write an SQL statement to give every employee a 10 percent raise.
6. Write an SQL statement to delete the Toy department. Given the referential integrity constraints you chose for this schema, explain what happens when this statement is executed.

Exercise 3.9 Consider the SQL query whose answer is shown in Figure 3.6.

1. Modify this query so that only the *login* column is included in the answer.
2. If the clause WHERE *S.gpa* >= 2 is added to the original query, what is the set of tuples in the answer?

Exercise 3.10 Explain why the addition of NOT NULL constraints to the SQL definition of the Manages relation (in Section 3.5.3) would not enforce the constraint that each department must have a manager. What, if anything, is achieved by requiring that the *ssn* field of Manages be non-null?

Exercise 3.11 Suppose that we have a ternary relationship R between entity sets A, B, and C such that A has a key constraint and total participation and B has a key constraint; these are the only constraints. A has attributes *a1* and *a2*, with *a1* being the key; B and C are similar. R has no descriptive attributes. Write SQL statements that create tables corresponding to this information so as to capture as many of the constraints as possible. If you cannot capture some constraint, explain why.

Exercise 3.12 Consider the scenario from Exercise 2.2, where you designed an ER diagram for a university database. Vrite SQL staternents to create the corresponding relations and capture as many of the constraints as possible. If you cannot: capture some constraints, explain why.

Exercise 3.13 Consider the university database from Exercise 2.3 and the ER diagram you designed. Write SQL statements to create the corresponding relations and capture as many of the constraints as possible. If you cannot capture some constraints, explain why.

Exercise 3.14 Consider the scenario from Exercise 2.4, where you designed an ER diagram for a company database. Write SQL statements to create the corresponding relations and capture as many of the constraints as possible. If you cannot capture some constraints, explain why.

Exercise 3.15 Consider the Notown database from Exercise 2.5. You have decided to recommend that Notown use a relational database system to store company data. Show the SQL statements for creating relations corresponding to the entity sets and relationship sets in your design. Identify any constraints in the ER diagram that you are unable to capture in the SQL statements and briefly explain why you could not express them.

Exercise 3.16 Translate your ER diagram from Exercise 2.6 into a relational schema, and show the SQL statements needed to create the relations, using only key and null constraints. If your translation cannot capture any constraints in the ER diagram, explain why.

In Exercise 2.6, you also modified the ER diagram to include the constraint that tests on a plane must be conducted by a technician who is an expert on that model. Can you modify the SQL statements defining the relations obtained by mapping the ER diagram to check this constraint?

Exercise 3.17 Consider the ER diagram that you designed for the Prescriptions-R-X chain of pharmacies in Exercise 2.7. Define relations corresponding to the entity sets and relationship sets in your design using SQL.

Exercise 3.18 Write SQL statements to create the corresponding relations to the ER diagram you designed for Exercise 2.8. If your translation cannot capture any constraints in the ER diagram, explain why.

Exercise 3.19 Briefly answer the following questions based on this schema:

Emp(e'id: integer, *ename:* string, *age:* integer, *salary:* real)
Works(*eid:* integer, *did:* integer, *pct_time:* integer)
Dept(did: integer, *budget:* real, *managerid:* integer)

1. Suppose you have a view SeniorEmp defined as follows:

```
CREATE VIEW SeniorEmp (sname, sage, salary)
       AS SELECT E.ename, Kage, E.salary
          FROM   Emp E
          WHERE  Kage > 50
```

   Explain what the system will do to process the following query:

```
SELECT S.sname
FROM   SeniorEmp S
WHERE  S.salary > 100,000
```

2. Give an example of a view on Emp that could be automatically updated by updating Emp.
3. Give an example of a view on Emp that would be impossible to update (automatically) and explain why your example presents the update problem that it does.

Exercise 3.20 Consider the following schema:

Suppliers(*sid:* integer, *sname:* string, *address:* string)
Parts(pid: integer, *pname:* string, *color:* string)
Catalog(sid: integer, *pid:* integer, *cost:* real)

The Catalog relation lists the prices charged for parts by Suppliers. Answer the following questions:

- Give an example of an updatable view involving one relation.
- Give an example of an updatable view involving two relations.
- Give an example of an insertable-into view that is updatable.
- Give an example of an insertable-into view that is not updatable.

PROJECT-BASED EXERCISES

Exercise **3.21** Create the relations Students, Faculty, Courses, Rooms, Enrolled, Teaches, and Meets_In in Minibase.

Exercise **3.22** Insert the tuples shown in Figures 3.1 and 3.4 into the relations Students and Enrolled. Create reasonable instances of the other relations.

Exercise **3.23** What integrity constraints are enforced by Minibase?

Exercise **3.24** Run the SQL queries presented in this chapter.

BIBLIOGRAPHIC NOTES

The relational model was proposed in a seminal paper by Codd [187]. Childs [176] and Kuhns [454] foreshadowed some of these developments. Gallaire and Minker's book [296] contains several papers on the use of logic in the context of relational databases. A system based on a variation of the relational model in which the entire database is regarded abstractly as a single relation, called the *universal relation,* is described in [746]. Extensions of the relational model to incorporate *null* values, which indicate an unknown or missing field value, are discussed by several authors; for example, [329, 396, 622, 754, 790].

Pioneering projects include System R [40, 150] at IBM San Jose Research Laboratory (now IBM Almaden Research Center), Ingres [717] at the University of California at Berkeley, PRTV [737] at the IBM UK Scientific Center in Peterlee, and QBE [801] at IBM T. J. Watson Research Center.

A rich theory underpins the field of relational databases. Texts devoted to theoretical aspects include those by··Atzeni and DeAntonellis [45]; Maier [501]; and Abiteboul, Hull, and Vianu [3]. [415] is an excellent survey article.

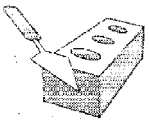
Integrity constraints in relational databases have been discussed at length. [190] addresses semantic extensions to the relational model, and integrity, in particular referential integrity. [360] discusses semantic integrity constraints. [203] contains papers that address various aspects of integrity constraints, including in particular a detailed discussion of referential integrity. A vast literature deals \vith enforcing integrity constraints. [51] compares the cost

of enforcing integrity constraints via compile-time, run-time, and post-execution checks. [145] presents an SQL-based language for specifying integrity constraints and identifies conditions under which integrity rules specified in this language can be violated. [713] discusses the technique of integrity constraint checking by query modification. [180] discusses real-time integrity constraints. Other papers on checking integrity constraints in databases include [82, 122, 138,517]. [681] considers the approach of verifying the correctness of programs that access the database instead of run-time checks. Note that this list of references is far fTom complete; in fact, it does not include any of the many papers on checking recursively specified integrity constraints. Some early papers in this widely studied area can be found in [296] and [295].

For references on SQL, see the bibliographic notes for Chapter 5. This book does not discuss specific products based on the relational model, but many fine books discuss each of the major commercial systems; for example, Chamberlin's book on DB2 [149], Date and McGoveran's book on Sybase [206], and Koch and Loney's book on Oracle [443].

Several papers consider the problem of translating updates specified on views into updates on the underlying table [59, 208, 422, 468, 778]. [292] is a good survey on this topic. See the bibliographic notes for Chapter 25 for references to work querying views and maintaining materialized views.

[731] discusses a design methodology based on developing an ER diagram and then translating to the relational model. Markowitz considers referential integrity in the context of ER to relational mapping and discusses the support provided in some commercial systems (as of that date) in [513, 514].



4

RELATIONAL ALGEBRA AND CALCULUS

- What is the foundation for relational query languages like SQL? What is the difference between procedural and declarative languages?
- What is relational algebra, and why is it important?
- What are the basic algebra operators, and how are they combined to write complex queries?
- What is relational calculus, and why is it important?
- What subset of mathematical logic is used in relational calculus, and how is it used to write queries?
- Key concepts: relational algebra, select, project, union, intersection, cross-product, join, division; tuple relational calculus, domain relational calculus, formulas, universal and existential quantifiers, bound and free variables

> Stand finn in your refusal to remain conscious during algebra. In real life, I assure you, there is no such thing as algebra.
>
> —Fran Lebowitz, *Social Studies*

This chapter presents two formal query languages associated with the relational model. Query languages are specialized languages for asking questions, or queries, that involve the data in a database. After covering some preliminaries in Section 4.1, we discuss *relational algebra* in Section 4.2. Queries in relational algebra are composed using a collection of operators, and each query describes a step-by-step procedure for computing the desired answer; that is, queries are

specified in an *operational* manner. In Section 4.3, we discuss *relational calculus*, in which a query describes the desired ans\ver without specifying how the answer is to be computed; this nonprocedural style of querying is called *declarat'i'Ve*. We usually refer to relational algebra and relational calculus as algebra and calculus, respectively. We compare the expressive power of algebra and calculus in Section 4.4. These formal query languages have greatly influenced commercial query languages such as SQL, which we discuss in later chapters.

4.1 PRELIMINARIES

We begin by clarifying some important points about relational queries. The inputs and outputs of a query are relations. A query is evaluated using *instances* of each input relation and it produces an instance of the output relation. In Section 3.4, we used field names to refer to fields because this notation makes queries more readable. An alternative is to always list the fields of a given relation in the same order and refer to fields by position rather than by field name.

In defining relational algebra and calculus, the alternative of referring to fields by position is more convenient than referring to fields by name: Queries often involve the computation of intermediate results, which are themselves relation instances; and if we use field names to refer to fields, the definition of query language constructs must specify the names of fields for all intermediate relation instances. This can be tedious and is really a secondary issue, because we can refer to fields by position anyway. On the other hand, field names make queries more readable.

Due to these considerations, we use the positional notation to formally define relational algebra and calculus. We also introduce simple conventions that allow intermediate relations to 'inherit' field names, for convenience.

We present a number of sample queries using the following schema:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)
Boats(*bid:* integer, *bnarne:* string, *coloT:* string)
Reserves(*sid:* integer, *bid:* integer, *day:* date)

The key fields are underlined, and the domain of each field is listed after the field name. Thus, *sid* is the key for Sailors, *bid* is the key for Boats, and all three fields together form the key for Reserves. Fields in an instance of one of these relations are referred to by name, or positionally, using the order in which they were just listed.

In several examples illustrating the relational algebra operators, we use the instances 81 and 82 (of Sailors) and *R1* (of Reserves) shown in Figures 4.1, 4.2, and 4.3, respectively.

sid			
22	Dustin	7	45.0
31	Lubber	8	55.5
58	Rusty	10	35.0

Figure 4.1 Instance *S1* of Sailors

sid		rating	age
28	yuppy	9	35.0
31	Lubber	8	55.5
44	guppy	5	35.0
58	Rusty	10	35.0

Figure 4.2 Instance *S2* of Sailors

sid	*bid*	*day*
22	101	10/10/96
58	103	11/12/96

Figure 4.3 Instance *R1* of Reserves

4.2 RELATIONAL ALGEBRA

Relational algebra is one of the two formal query languages associated with the relational model. Queries in algebra are composed using a collection of operators. A fundamental property is that every operator in the algebra accepts (one or two) relation instances as arguments and returns a relation instance as the result. This property makes it easy to *compose* operators to form a complex query-a relational algebra expression is recursively defined to be a relation, a unary algebra operator applied to a single expression, or a binary algebra operator applied to two expressions. We describe the basic operators of the algebra (selection, projection, union, cross-product, and difference), as well as some additional operators that can be defined in terms of the basic operators but arise frequently enough to warrant special attention, in the following sections.

Each relational query describes a step-by-step procedure for computing the desired answer, based on the order in which operators are applied in the query. The procedural nature of the algebra allows us to think of an algebra expression as a recipe, or a plan, for evaluating a query, and relational systems in fact use algebra expressions to represent query evaluation plans.

4.2.1 Selection and Projection

Relational algebra includes operators to *select* rows from a relation (σ) and to *project* columns (π). These operations allow us to manipulate data in a single relation. Consider the instance of the Sailors relation shown in Figure 4.2, denoted as *52.* We can retrieve rows corresponding to expert sailors by using the σ operator. The expression

$$\sigma_{rating>8}(52)$$

evaluates to the relation shown in Figure 4.4. The subscript *rating*> 8 specifies the selection criterion to be applied while retrieving tuples.

sid	*sname*	*rating*	*age*
28	yuppy	9	35.0
58	Rusty	10	35.0

Figure 4.4 $\sigma_{rating>8}(S2)$

sname	*rating*
yuppy	9
Lubber	8
guppy	5
Rusty	10

Figure 4.5 $\pi_{sname,rating}(S2)$

The selection operator σ specifies the tuples to retain through a *selection condition.* In general, the selection condition is a Boolean combination (i.e., an expression using the logical connectives /\ and V) of *terms* that have the form *attribute* op *constant* or *attribute1* op *attribute2*, where op is one of the comparison operators <, <=, =, $\neq$, >=, or >. The reference to an attribute can be by position (of the form .i or i) or by name (of the form *.name* or *name).* The schema of the result of a selection is the schema of the input relation instance.

The projection operator π allows us to extract columns from a relation; for example, we can find out all sailor names and ratings by using 1f. The expression

$$\pi_{sname,rafing(52)}$$

evaluates to the relation shown in Figure 4.5. The subscript *8na:me)rating* specifies the fields to be retained; the other fields are 'projected out.' The schema of the result of a projection is determined by the fields that are projected in the obvious way.

Suppose that we wanted to find out only the ages of sailors. The expression

$$\pi_{age}(S2)$$

evaluates to the relation shown in Figure 4.6. The irnportant point to note is that, although three sailors are aged 35, a single tuple with *age=35.0* appears in

the result of the projection. This follm\'8 from the definition of a relation as a *set* of tuples. In practice, real systems often omit the expensive step of eliminating *duplicate tuples*, leading to relations that are multisets. However, our discussion of relational algebra and calculus assumes that duplicate elimination is always done so that relations are always sets of tuples.

Since the result of a relational algebra expression is always a relation, we can substitute an expression wherever a relation is expected. For example, we can compute the names and ratings of highly rated sailors by combining two of the preceding queries. The expression

$$\pi_{sname,rating}(\sigma_{rating>8}(S2))$$

produces the result shown in Figure 4.7. It is obtained by applying the selection to S2 (to get the relation shown in Figure 4.4) and then applying the projection.

age
35.0
55.5

Figure 4.6 $\pi_{age}(S2)$

sname	*rating*
yuppy	9
Rusty	10

Figure 4.7 $\pi_{sname,rating}(\sigma_{rating>8}(S2))$

4.2.2 Set Operations

The following standard operations on sets are also available in relational algebra: *un'ion* (U), *intersection* (n), *set-difference* (-), and *cross-product* (x).

- **Union:** $R \cup S$ returns a relation instance containing aU tuples that occur in *either* relation instance R or relation instance S (or both). Rand S must be *union-compatible,* and the schema of the result is defined to be identical to the schema of R.

  Two relation instances are said to be **union-compatible** if the following conditions hold:
  - they have the same number of the fields, and
  - corresponding fields, taken in order from left to right, have the same *domains.*

  Note that field names are not used in defining union-compatibility. For convenience, we will assume that the fields of $R \cup S$ inherit names from R, if the fields of R have names. (This assumption is implicit in defining the schema of $R \cup S$ to be identical to the schema of R, as stated earlier.)

- **Intersection:** $R \cap S$ returns a relation instance containing all tuples that occur in *both* Rand S. The relations Rand S must be union-compatible, and the schema of the result is defined to be identical to the schema of R.

- Set-difference: $R-S$ returns a relation instance containing all tuples that occur in R but not in S. The relations R and S must be union-compatible, and the schema of the result is defined to be identical to the schema of R.
- Cross-product: $R \times S$ returns a relation instance whose schema contains all the fields of R (in the same order as they appear in R) followed by all the fields of S (in the same order as they appear in S). The result of $R \times S$ contains one tuple (r, s) (the concatenation of tuples r and s) for each pair of tuples $r \in R$, $s \in S$. The cross-product opertion is sometimes called Cartesian product.

  We use the convention that the fields of $R \times S$ inherit names from the corresponding fields of R and S. It is possible for both R and S to contain one or more fields having the same name; this situation creates a *naming conflict*. The corresponding fields in $R \times S$ are unnamed and are referred to solely by position.

In the preceding definitions, note that each operator can be applied to relation instances that are computed using a relational algebra (sub)expression.

We now illustrate these definitions through several examples. The union of $S1$ and $S2$ is shown in Figure 4.8. Fields are listed in order; field names are also inherited from $S1$. $S2$ has the same field names, of course, since it is also an instance of Sailors. In general, fields of $S2$ may have different names; recall that we require only domains to match. Note that the result is a *set* of tuples. Tuples that appear in both $S1$ and $S2$ appear only once in $S1 \cup S2$. Also, $S1 \cup R1$ is not a valid operation because the two relations are not union-compatible. The intersection of $S1$ and $S2$ is shown in Figure 4.9, and the set-difference $S1 - S2$ is shown in Figure 4.10.

sid	*sname*	*rating*	*age*
22	Dustin	7	45.0
31	Lubber	8	55.5
58	Rusty	10	35.0
28	yuppy	9	35.0
44	guppy	5	35.0

Figure 4.8 $S1 \cup S2$

The result of the cross-product $S1 \times R1$ is shown in Figure 4.11. Because $R1$ and $S1$ both have a field named *sid*, by our convention on field names, the corresponding two fields in $S1 \times R1$ are unnamed, and referred to solely by the position in which they appear in Figure 4.11. The fields in $S1 \times R1$ have the same domains as the corresponding fields in $R1$ and $S1$. In Figure 4.11, *sid* is

sid	sname		
31	Lubber	8	55.5
58	Rusty	10	35.0

Figure 4.9 S1 ∩ S2

sid	sname	rating	age
22	Dustin	7	45.0

Figure 4.10 S1 - S2

listed in parentheses to emphasize that it is not an inherited field name; only the corresponding domain is inherited.

(sid)	sname	rating	age	(sid)	bid	day
22	Dustin	7	45.0	22	101	10/10/96
22	Dustin	7	45.0	58	103	11/12/96
31	Lubber	8	55.5	22	101	10/10/96
31	Lubber	8	55.5	58	103	11/12/96
58	Rusty	10	35.0	22	101	10/10/96
58	Rusty	10	35.0	58	103	11/12/96

Figure 4.11 S1 x *R1*

4.2.3 Renaming

We have been careful to adopt field name conventions that ensure that the result of a relational algebra expression inherits field names from its argument (input) relation instances in a natural way whenever possible. However, name conflicts can arise in some cases; for example, in S1 x R1. It is therefore convenient to be able to give names explicitly to the fields of a relation instance that is defined by a relational algebra expression. In fact, it is often convenient to give the instance itself a name so that we can break a large algebra expression into smaller pieces by giving names to the results of subexpressions.

We introduce a renaming operator ρ for this purpose. The expression $\rho(R(F), E)$ takes an arbitrary relational algebra expression E and returns an instance of a (new) relation called R. R contains the same tuples as the result of E and has the same schema as E, but some fields are renamed. The field names in relation R are the sarne as in E, except for fields renamed in the *renaming list* F, which is a list of terms having the form $oldname \rightarrow newname$ or $position \rightarrow newname$. For ρ to be well-defined, references to fields (in the form of *oldnames* or *posit.ions* in the renaming list) may be unarnbiguous and no two fields in the result may have the same name. Sometimes we want to only rename fields or (re)name the relation; we therefore treat both R and F as optional in the use of ρ. (Of course, it is meaningless to omit both.)

For example, the expression $p(C(1 \rightarrow s'id1,5 \rightarrow sid2),$ 81 x $R1)$ returns a relation that contains the tuples shown in Figure 4.11 and has the following schema: *C(sidl:* integer, *sname:* string, *rating:* integer, *age:* real, *sid2:* integer, *bid:* integer, *day:* dates).

It is customary to include some additional operators in the algebra, but all of them can be defined in terms of the operators we have defined thus far. (In fact, the renaming operator is needed only for syntactic convenience, and even the n operator is redundant; $R \cap 8$ can be defined as $R - (R - 8)$.) We consider these additional operators and their definition in terms of the basic operators in the next two subsections.

4.2.4 Joins

The *join* operation is one of the most useful operations in relational algebra and the most commonly used way to combine information from two or more relations. Although a join can be defined as a cross-product followed by selections and projections, joins arise much more frequently in practice than plain cross-products. Further, the result of a cross-product is typically much larger than the result of a join, and it is very important to recognize joins and implement them without materializing the underlying cross-product (by applying the selections and projections 'on-the-fly'). For these reasons, joins have received a lot of attention, and there are several variants of the join operation.[1]

Condition Joins

The most general version of the join operation accepts a *join condition* c and a pair of relation instances as arguments and returns a relation instance. The *join cond'it-ion* is identical to a *selection condition* in form. The operation is defined as follows:

$$R \bowtie_c S = \sigma_c(R \times S)$$

Thus $\bowtie$ is defined to be a cross-product followed by a selection. Note that the condition c can (and typically *does)* refer to attributes of both Rand *S*. The reference to an attribute of a relation, say, *R,* can be by positioll (of the form *R.i)* or by Ilame (of the form *R.name)*.

As an example, the result of $SI \bowtie_{S1.sid<R1.sid} R1$ is shown in Figure 4.12. Because *sid* appears in both 81 and *R1,* the corresponding fields in the result of the cross-product 81 x *R1* (and therefore in the result of $81 \bowtie_{S1.sid<R1.sid} R1$)

[1] Several variants of joins are not discussed in this chapter. An important class of joins, called *outer joins*, is discussed in Chapter 5.

are unnamed. Domains are inherited from the corresponding fields of 81 and Rl.

(sid)	sname	rating	age	(sid)	bid	day
22	Dustin	7	45.0	58	103	11/12/96
31	Lubber	8	55.5	58	103	11/12/96

Figure 4.12 51 *NSI.sid<R1.sid R1*

Equijoin

A common special case of the join operation $R \bowtie S$ is when the *join condition* consists solely of equalities (connected by *1\)* of the form *R.name1 = 8.name2,* that is, equalities between two fields in Rand *S.* In this case, obviously, there is some redundancy in retaining both attributes in the result. For join conditions that contain only such equalities, the join operation is refined by doing an additional projection in which *8.name2* is dropped. The join operation with this refinement is called **equijoin.**

The schema of the result of an equijoin contains the fields of *R* (with the same names and domains as in *R)* followed by the fields of *S* that do not appear in the join conditions. If this set of fields in the result relation includes two fields that inherit the same name from Rand 8, they are unnamed in the result relation.

We illustrate $S1 \bowtie_{R.sid=S.sid} Rl$ in Figure 4.13. Note that only one field called *sid* appears in the result.

sid	sname	rating	age	bid·1	day
22	Dustin	7	45.0	101	10/10/96
58	Rusty	10	35.0	103	11/12/96

Figure 4.13 81 $\bowtie_{R.sid=S.sid}$ *HI*

Natural **Join**

A further special case of the join operation $R \bowtie S$ is an eqUlJom in which equalities arc specified on *all* fields having the same name in Rand *S.* In this case, we can simply omit the join condition; the default is that the join condition is a collection of equalities on all common fields. We call this special case a *natural jo'in,* and it has the nice property that the result is guaranteed not to have two fields with the saIne name.

The equijoin expression $S1 \bowtie_{R.sid=S.sid} R1$ is actually a natural join and can simply be denoted as $S1 \bowtie R1$, since the only common field is *sid.* If the two relations have no attributes in common, $S1 \bowtie Rl$ is simply the cross-product.

4.2.5 Division

The division operator is useful for expressing certain kinds of queries for example, "Find the names of sailors who have reserved all boats." Understanding how to use the basic operators of the algebra to define division is a useful exercise. However, the division operator does not have the same importance as the other operators-it is not needed as often, and database systems do not try to exploit the semantics of division by implementing it as a distinct operator (as, for example, is done with the join operator).

We discuss division through an example. Consider two relation instances A and B in which A has (exactly) two fields x and y and B has just one field y, with the same domain as in A. We define the *division* operation A/B as the set of all x values (in the form of unary tuples) such that for *every* y value in (a tuple of) B, there is a tuple (x,y) in A.

Another way to understand division is as follows. For each x value in (the first column of) A, consider the set of y values that appear in (the second field of) tuples of A with that x value. If this set contains (all y values in) B, the x value is in the result of A/B.

An analogy with integer division may also help to understand division. For integers A and B, A/B is the largest integer Q such that $Q * B \leq A$. :For relation instances A and B, A/B is the largest relation instance Q such that $Q \times B \subseteq A$.

Division is illustrated in Figure 4.14. It helps to think of A as a relation listing the parts supplied by suppliers and of the B relations as listing parts. A/Bi computes suppliers who supply *all* parts listed in relation instance Bi.

Expressing A/Bin terms of the basic algebra operators is an interesting exercise, and the reader should try to do this before reading further. The basic idea is to compute all x values in A that are not *disqualified.* An x value is *disqualified* if by attaching a y value from B, we obtain a tuple (x,y) that is not in A. We can compute disqualified tuples using the algebra expression

$$\pi_x((\pi_x(A) \times B) - A)$$

Thus, we can define A/B as

$$\pi_x(A) - \pi_x((\pi_x(A) \times B) - A)$$

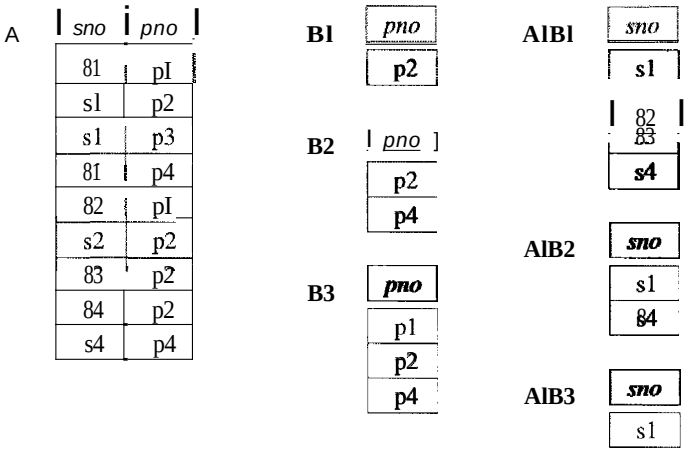


Figure 4.14 Examples Illustrating Division

To understand the division operation in full generality, we have to consider the case when both *x* and yare replaced by a set of attributes. The generalization is straightforward and left as an exercise for the reader. We discuss two additional examples illustrating division (Queries Q9 and Q10) later in this section.

4.2.6 More Examples of Algebra Queries

We now present several examples to illustrate how to write queries in relational algebra. We use the Sailors, Reserves, and Boats schema for all our examples in this section. We use parentheses as needed to make our algebra expressions unambiguous. Note that all the example queries in this chapter are given a unique query number. The query numbers are kept unique across both this chapter and the SQL query chapter (Chapter 5). This numbering makes it easy to identify a query when it is revisited in the context of relational calculus and SQL and to compare different ways of writing the same query. (All references to a query can be found in the subject index.)

In the rest of this chapter (and in Chapter 5), we illustrate queries using the instances 83 of Sailors, *R2* of Reserves, and *B1* of Boats, shown in Figures 4.15, 4.16, and 4.17, respectively.

(Q1) Find the names of sailors who have rcscT'ucd boat 103.

This query can be written as follows:

$$\pi_{sname}((\sigma_{bid=103} Reserves) \bowtie 8ailoT.5)$$

sid		rating	age
22	Dustin	7	45.0
29	Brutus	1	33.0
31	Lubber	8	55.5
32	Andy	8	25.5
58	Rusty	10	35.0
64	Horatio	7	35.0
71	Zorba	10	16.0
74	Horatio	9	35.0
85	Art	3	25.5
95	Bob	3	63.5

Figure 4.15 An Instance 83 of Sailors

sid	bid	day
22	101	10/10/98
22	102	10/10/98
22	103	10/8/98
22	104	10/7/98
31	102	11/10/98
31	103	11/6/98
31	104	11/12/98
64	101	9/5/98
64	102	9/8/98
74	103	9/8/98

Figure 4.16 An Instance *R2* of Reserves

We first compute the set of tuples in Reserves with *bid* = 103 and then take the natural join of this set with Sailors. This expression can be evaluated on instances of Reserves and Sailors. Evaluated on the instances *R2* and *S3*, it yields a relation that contains just one field, called *sname*, and three tuples *(Dustin)*, *(Horatio)*, and *(Lubber)*. (Observe that two sailors are called Horatio and only one of them has reserved a red boat.)

bid	bname	color
101	Interlake	blue
102	Interlake	red
103	Clipper	green
104	Marine	red

Figure 4.17 An Instance *HI* of Boats

We can break this query into smaller pieces llsing the renaming operator *p:*

$$p(Temp1, \sigma_{bid=103} ReseTves)$$
$$p(Temp2, Temp1 \bowtie Sailor's)$$
$$\pi_{sname}(Temp2)$$

Notice that because we are only llsing *p* to give names to intermediate relations, the renaming list is optional and is omitted. *TempI* denotes an intermediate relation that identifies reservations of boat 103. *Temp2* is another intermediate relation, and it denotes sailors who have made a reservation in the set *Templ*. The instances of these relations when evaluating this query on the instances *R2* and *S3* are illustrated in Figures 4.18 and 4.19. Finally, we extract the *sname* column from *Temp2*.

sid	bid	day
22	103	10/8/98
31	103	11/6/98
74	103	9/8/98

Figure 4.18 Instance of *TempI*

sid	sname	rating	age	bid	
22	Dustin	7	45.0	103	10/8/98
31	Lubber	8	55.5	103	11/6/98--
74	Horatio	9	35.0	103	9/8/98

Figure 4.19 Instance of *Temp2*

The version of the query using p is essentially the same as the original query; the use of p is just syntactic sugar. However, there are indeed several distinct ways to write a query in relational algebra. Here is another way to write this query:

$$Jrsname(CJbid=103(Reserves \text{ IX! } Sailors))$$

In this version we first compute the natural join of Reserves and Sailors and then apply the selection and the projection.

This example offers a glimpse of the role played by algebra in a relational DBMS. Queries are expressed by users in a language such as SQL. The DBMS translates an SQL query into (an extended form of) relational algebra and then looks for other algebra expressions that produce the same answers but are cheaper to evaluate. If the user's query is first translated into the expression

$$\pi_{sname}(CJbid=103\,(Reserves \text{ IX! } Sailors))$$

a good query optimizer will find the equivalent expression

$$\pi sname\,((CJb{\cdot}id=103Reserves) \text{ IX! } Sailors)$$

Further, the optimizer will recognize that the second expression is likely to be less expensive to compute because the sizes of intermediate relations are smaller, thanks to the early use of selection.

(Q2) Find the names of sailors who ha've reserved a red boat.

$$\pi_{sname}((\sigma_{color='red'}Boats) \text{ IX! } Reserves \bowtie SailoI's)$$

This query involves a series of two joins. First, we choose (tuples describing) red boats. Then, we join this set with Reserves (natural join, with equality specified on the *bid* column) to identify reservations of red boats. Next, we join the resulting intermediate relation with Sailors (natural join, with equality specified on the *sid* column) to retrieve the names of sailors who have rnade reservations for red boats. Finally, we project the sailors' names. The answer, when evaluated on the instances *B1*, *R2*, and *S3*, contains the names Dustin, Horatio, and Lubber.

An equivalent expression is:

$$\pi_{sname}(\pi_{sid}((\pi_{bid}\sigma_{color='red'}Boats) \bowtie Reserves) \bowtie Sailors)$$

The reader is invited to rewrite both of these queries by using *p* to make the intermediate relations explicit and compare the schemas of the intermediate relations. The second expression generates intermediate relations with fewer fields (and is therefore likely to result in intermediate relation instances with fewer tuples as well). A relational query optimizer would try to arrive at the second expression if it is given the first.

(Q3) Find the colors of boats reserved by Lubber.

$$\pi_{color}((\sigma_{sname='Lubber'}Sailors) \bowtie Reserves \bowtie Boats)$$

This query is very similar to the query we used to compute sailors who reserved red boats. On instances *Bl, R2,* and *S3,* the query returns the colors green and red.

(Q4) Find the names of sailors who have reserved at least one boat.

$$Jrsname(Sailors \bowtie Reserves)$$

The join of Sailors and Reserves creates an intermediate relation in which tuples consist of a Sailors tuple 'attached to' a Reserves tuple. A Sailors tuple appears in (some tuple of) this intermediate relation only if at least one Reserves tuple has the same *sid* value, that is, the sailor has made some reservation. The answer, when evaluated on the instances *Bl, R2* and *S3,* contains the three tuples $\langle Dustin \rangle$, $\langle HoTatio \rangle$, and $\langle Lubber \rangle$. Even though two sailors called Horatio have reserved a boat, the answer contains only one copy of the tuple $\langle HoTatio \rangle$, because the answer is a *relation,* that is, a *set* of tuples, with no duplicates.

At this point it is worth remarking on how frequently the natural join operation is used in our examples. This frequency is more than just a coincidence based on the set of queries we have chosen to discuss; the natural join is a very natural, widely used operation. In particular, natural join is frequently used when joining two tables on a foreign key field. In Query Q4, for exalnple, the join equates the *sid* fields of Sailors and Reserves, and the *sid* field of Reserves is a foreign key that refers to the *sid* field of Sailors.

(Q5) Find the narnes of sailors who have reserved a Ted 0T a gTeen boat.

$$p(Tempboats, (acoloT='rcd' Boats) \cup (\sigma_{color='green'}Boats))$$
$$\pi_{sname}(Tempboats \bowtie ReseTves \bowtie Sailors)$$

We identify the set of all boats that are either red or green (Tempboats, which contains boats \vith the *bid*s 102, 103, and 104 on instances *E1, R2,* and *S3).* Then we join with Reserves to identify *sid*s of sailors who have reserved Olle of these boats; this gives us *sid*s 22, 31, 64, and 74 over our example instances. Finally, we join (an intermediate relation containing this set of *sids)* with Sailors to find the names of Sailors with these *sid*s. This gives us the names Dustin, Horatio, and Lubber on the instances $B1$, *R2,* and *S3.* Another equivalent definition is the following:

p(Tempboats, (acolor='red'Vcolor='green' Boats))
7fsname(Tempboats ⋈ *Reserves* ⋈ *Sailors)*

Let us now consider a very similar query.

(Q6) Find the names of sailors who have reserved a red and a green boat. It is tempting to try to do this by simply replacing U by n in the definition of Tempboats:

p(Tempboats2, (acolor='red,Eoats) n *(O"color='green,Boats))*
$\pi_{sname}(Tempboats2 \bowtie Reserves \bowtie Sailors)$

However, this solution is incorrect-it instead tries to compute sailors who have reserved a boat that is both red and green. (Since *bid* is a key for Boats, a boat can be only one color; this query will always return an empty answer set.) The correct approach is to find sailors who have reserved a red boat, then sailors who have reserved a green boat, and then take the intersection of these two sets:

p(Tempred, π_{sid}*((acolor='red' Eoats)* ⋈ *Reserves))*
p(Tempgreen, $\pi_{sid}((\sigma_{color='green'} Boats) \bowtie Reserves))$
$\pi_{sname}((Tempred$ n *Tempgreen)* ⋈ *Sailors)*

The two temporary relations compute the *sid*s of sailors, and their intersection identifies sailors who have reserved both red and green boats. On instances $B1$, *R2,* and *53,* the *sid*s of sailors who have reserved a red boat are 22, 31, and 64. The *sid*s of sailors who have reserved a green boat are 22, 31, and 74. Thus, sailors 22 and 31 have reserved both a red boat and a green boat; their names are Dustin and Lubber.

This formulation of Query Q6 can easily be adapted to find sailors who have reserved red or green boats (Query Q5); just replace n by U:

$\rho(Tempred, \pi_{sid}((\sigma_{color='red'} Boats) \bowtie Reserves))$
p(Tempgreen, π_{sid}*((O"color='green' Boats)* ⋈ *Reserves))*
$\pi_{sname}((Tempred$ U *Tempgreen)* ⋈ *Sailors)*

In the formulations of Queries Q5 and Q6, the fact that *sid* (the field over which we compute union or intersection) is a key for Sailors is very important. Consider the following attempt to answer Query Q6:

$$p(Tempred, Jrsname((CJcolor='red,Boats) \bowtie Reserves \bowtie Sailors))$$
$$p(Tempgreen,Jrsname((CJcoloT='gTeenlBoats) \bowtie Reserves \bowtie Sailors))$$
$$Tempred \cap Tempgreen$$

This attempt is incorrect for a rather subtle reason. Two distinct sailors with the same name, such as Horatio in our example instances, may have reserved red and green boats, respectively. In this case, the name Horatio (incorrectly) is included in the answer even though no one individual called Horatio has reserved a red boat and a green boat. The cause of this error is that *sname* is used to identify sailors (while doing the intersection) in this version of the query, but *sname* is not a key.

(Q7) Find the names of sailors who have reser-ved at least two boats.

$$p(Reser\text{-}vations, \pi_{sid,sname,bid}(Sailors \bowtie Reserves))$$
$$p(Reservationpairs(1 \rightarrow sid1, 2 \rightarrow sname1, 3 \rightarrow bid1, 4 \rightarrow sid2,$$
$$5 \rightarrow sname2, 6 \rightarrow bid2), Reservations \times Reservations)$$
$$\pi_{sname1} CJ_{(sid1=sid2)\wedge(bid1\neq bid2)} Reservationpair\text{-}s$$

First, we compute tuples of the form *(sid,sname, bid)*, where sailor *sid* has made a reservation for boat *bid;* this set of tuples is the temporary relation Reservations. Next we find all pairs of Reservations tuples where the same sailor has made both reservations and the boats involved are distinct. Here is the central idea: To show that a sailor has reserved two boats, we must find two Reservations tuples involving the same sailor but distinct boats. Over instances $B1$, *R2,* and *S3,* each of the sailors with *sids* 22, 31, and 64 have reserved at least two boats. Finally, we project the names of such sailors to obtain the answer, containing the names Dustin, Horatio, and Lubber.

Notice that we included *sid* in Reservations because it is the key field identifying sailors, and we need it to check that two Reservations tuples involve the same sailor. As noted in the previous example, we cannot use *sname* for this purpose.

(Q8) Find the sids of sailors with age over 20 *who have not TeseTved a Ted boat.*

$$\pi_{sid}(\sigma_{age>20} Sailors) -$$
$$\tau_{rsid}((CJ_{C0}[oT='red,Boats) \bowtie Reserves \bowtie Sailors)$$

This query illustrates the use of the set-difference operator. Again, we use the fact that *sid* is the key for Sailors. We first identify sailors aged over 20 (over

instances $B1$, $R2$, and $S3$, sids 22, 29, 31, 32, 58, 64, 74, 85, and 95) and then discard those who have reserved a red boat (sids 22, 31, and 64), to obtain the answer (sids 29, 32, 58, 74, 85, and 95). If we want to compute the names of such sailors, we must first compute their sids (as shown earlier) and then join with Sailors and project the *sname* values.

(Q9) Find the names of sailors who have reserved all boats.

The use of the word *all* (or *every*) is a good indication that the division operation might be applicable:

$$p(Tempsids, (\pi_{sid,bid}Reserves)/(\pi_{bid}Boats))$$
$$\pi_{sname}(Tempsids \bowtie Sailors)$$

The intermediate relation Tempsids is defined using division and computes the set of *sids* of sailors who have reserved every boat (over instances *B1*, *R2*, and *S3*, this is just *sid* 22). Note how we define the two relations that the division operator (/) is applied to---the first relation has the schema *(sid,bid)* and the second has the schema *(bid)*. Division then returns all *sids* such that there is a tuple *(sid,bid)* in the first relation for each *bid* in the second. Joining Tempsids with Sailors is necessary to associate names with the selected *sids;* for sailor 22, the name is Dustin.

(Q10) Find the names of sailors who have reserved all boats called Interlake.

$$p(Tempsids, (\pi_{sid,bid}Reserves)/(\pi_{bid}(\sigma_{bname='Interlake'}Boats)))$$
$$\pi_{sname}(Tempsids \bowtie Sailors)$$

The only difference with respect to the previous query is that now we apply a selection to Boats, to ensure that we compute *bids* only of boats named *Interlake* in defining the second argument to the division operator. Over instances *E1*, *R2*, and *S3*, Tempsids evaluates to *sids* 22 and 64, and the answer contains their names, Dustin and Horatio.

4.3 RELATIONAL CALCULUS

Relational calculus is an alternative to relational algebra. In contrast to the algebra, which is procedural, the calculus is nonprocedural, or *declarative*, in that it allows us to describe the set of answers without being explicit about how they should be computed. Relational calculus has had a big influence on the design of commercial query languages such as SQL and, especially, Query-by-Example (QBE).

The variant of the calculus we present in detail is called the **tuple relational calculus** (TRC). Variables in TRC take on tuples as values. In another vari-

ant, called the domain relational calculus (DRC), the variables range over field values. TRC has had more of an influence on SQL, while DRC has strongly influenced QBE. We discuss DRC in Section 4.3.2.[2]

4.3.1 Tuple Relational Calculus

A tuple variable is a variable that takes on tuples of a particular relation schema as values. That is, every value assigned to a given tuple variable has the same number and type of fields. A tuple relational calculus query has the form $\{ T \mid p(T) \}$, where T is a tuple variable and $p(T)$ denotes a *formula* that describes T; we will shortly define formulas and queries rigorously. The result of this query is the set of all tuples t for which the formula $p(T)$ evaluates to true with $T = t$. The language for writing formulas $p(T)$ is thus at the heart of TRC and essentially a simple subset of *first-order logic*. As a simple example, consider the following query.

(Q11) Find all sailors with a rating above 7.

$$\{S \mid S \in Sailors \wedge S.rating > 7\}$$

When this query is evaluated on an instance of the Sailors relation, the tuple variable S is instantiated successively with each tuple, and the test $S.rating > 7$ is applied. The answer contains those instances of S that pass this test. On instance *S3* of Sailors, the answer contains Sailors tuples with *sid* 31, 32, 58, 71, and 74.

Syntax of TRC Queries

We now define these concepts formally, beginning with the notion of a formula. Let *Rel* be a relation name, R and S be tuple variables, a be an attribute of R, and b be an attribute of S. Let op denote an operator in the set $\{<, >, =, \leq, \geq, \neq\}$. An atomic formula is one of the following:

- $R \in Rel$
- $R.a$ op $S.b$
- $R.a$ op *constant*, or *constant* op $R.a$

A formula is recursively defined to be one of the following, where p and q are themselves formulas and $p(R)$ denotes a formula in which the variable R appears:

2The material on DRC is referred to in the (online) chapter on QBE; with the exception of this chapter, the material on DRC and TRC can be omitted without loss of continuity.

- any atomic formula
- $\neg p$, $P \wedge q$, $P \vee q$, or $p \Rightarrow q$
- $\exists R(p(R))$, where R is a tuple variable
- $\forall R(p(R))$, where R is a tuple variable

In the last two clauses, the quantifiers $\exists$ and $\forall$ are said to bind the variable R. A variable is said to be free in a formula or *subformuia* (a formula contained in a larger formula) if the (sub)formula does not contain an occurrence of a quantifier that binds it.[3]

We observe that every variable in a TRC formula appears in a subformula that is atomic, and every relation schema specifies a domain for each field; this observation ensures that each variable in a TRC formula has a well-defined domain from which values for the variable are drawn. That is, each variable has a well-defined *type,* in the programming language sense. Informally, an atomic formula $R \in Rel$ gives R the type of tuples in *Rel,* and comparisons such as *R.a* op *S.b* and *R.a* op *constant* induce type restrictions on the field *R.a*. If a variable R does not appear in an atomic formula of the form $R \in Rel$ (Le., it appears only in atomic formulas that are comparisons), we follow the convention that the type of R is a tuple whose fields include all (and only) fields of R that appear in the formula.

We do not define types of variables formally, but the type of a variable should be clear in most cases, and the important point to note is that comparisons of values having different types should always fail. (In discussions of relational calculus, the simplifying assumption is often made that there is a single domain of constants and this is the domain associated with each field of each relation.)

A TRC query is defined to be expression of the form $\{T \mid p(T)\}$, where T is the only free variable in the formula p.

Semantics of TRC Queries

What does a TRC query mean? More precisely, what is the set of answer tuples for a given TRC query? The answer to a TRC query $\{T \mid p(T)\}$, as noted earlier, is the set of all tuples t for which the formula *peT)* evaluates to true with variable T assigned the tuple value t. To complete this definition, we must state which assignments of tuple values to the free variables in a formula make the formula evaluate to true.

[3]We make the assumption that each variable in a formula is either free or bound by exactly one occurrence of a quantifier, to avoid worrying about details such as nested occurrences of quantifiers that bind some, but not all, occurrences of variables.

A query is evaluated on a given instance of the database. Let each free variable in a formula F be bound to a tuple value. For the given assignment of tuples to variables, with respect to the given database instance, F evaluates to (or simply 'is') true if one of the following holds:

- F is an atomic formula $R \in Rel$, and R is assigned a tuple in the instance of relation *Rel*.
- F is a comparison $R.a$ op $S.b$, $R.a$ op *constant*, or *constant* op $R.a$, and the tuples assigned to Rand S have field values $R.a$ and $S.b$ that make the comparison true.
- F is of the form $\neg p$ and p is not true, or of the form $p \wedge q$, and both p and q are true, or of the form $p \vee q$ and one of them is true, or of the form $p \Rightarrow q$ and q is true whenever[4] p is true.
- F is of the form $\exists R(p(R))$, and there is some assignment of tuples to the free variables in $p(R)$, including the variable R,[5] that makes the formula $p(R)$ true.
- F is of the form $\forall R(p(R))$, and there is some assignment of tuples to the free variables in $p(R)$ that makes the formula $p(R)$ true no matter what tuple is assigned to R.

Examples of TRC Queries

We now illustrate the calculus through several examples, using the instances *B1* of Boats, *R2* of Reserves, and *S3* of Sailors shown in Figures 4.15, 4.16, and 4.17. We use parentheses as needed to make our formulas unambiguous. Often, a formula $p(R)$ includes a condition $R \in Rel$, and the meaning of the phrases *some tuple R* and *for all tuples R* is intuitive. We use the notation $\exists R \in Rel(p(R))$ for $\exists R(R \in Rel \wedge p(R))$. Similarly, we use the notation $\forall R \in Rel(p(R))$ for $\forall R(R \in Rel \Rightarrow p(R))$.

(Q12) *Find the names and ages of sailors with a rating above 7.*

$$\{P \mid \exists S \in Sailors(S.rating > 7 \wedge P.name = S.sname \wedge P.age = S.age)\}$$

This query illustrates a useful convention: P is considered to be a tuple variable with exactly two fields, which are called *name* and *age*, because these are the only fields of P mentioned and P does not range over any of the relations in the query; that is, there is no subformula of the form $P \in Relname$. The result of this query is a relation with two fields, *name* and *age*. The atomic

4 *WheneveT* should be read more precisely as 'for all assignments of tuples to the free variables.'
5Note that some of the free variables in $p(R)$ (e.g., the variable R itself) IIay be bound in P.

formulas *P.name* = *S.sname* and *Page* = *S.age* give values to the fields of an answer tuple *P*. On instances *E1*, *R2*, and *S3*, the answer is the set of tuples $\langle Lubber, 55.5\rangle$, $\langle Andy, 25.5\rangle$, $\langle Rusty, 35.0\rangle$, *(Zorba,* 16.0), and $\langle Horatio, 35.0\rangle$.

(Q1S) Find the sailor name, boat'id, and reservation date for each reservation.

$\{P \mid \exists R \in Reserves\ \exists S \in Sailors$
$(R.sid = 8.sid \wedge P.bid = R.bid \wedge P.day = R.day \wedge P.sname = S.sname)\}$

For each Reserves tuple, we look for a tuple in Sailors with the same *sid*. Given a pair of such tuples, we construct an answer tuple *P* with fields *sname*, *bid*, and *day* by copying the corresponding fields from these two tuples. This query illustrates how we can combine values from different relations in each answer tuple. The answer to this query on instances *E1*, *R2*, and 83 is shown in Figure 4.20.

sname	*bid*	*day*
Dustin	101	10/10/98
Dustin	102	10/10/98
Dustin	103	10/8/98
Dustin	104	10/7/98
Lubber	102	11/10/98
Lubber	103	11/6/98
Lubber	104	11/12/98
Horatio	101	9/5/98
Horatio	102	9/8/98
Horatio	103	9/8/98

Figure 4.20 Answer to Query Q13

(Q1) Find the names of sailors who have reserved boat 103.

$\{P \mid \exists S \in Sailors\ \exists R \in Reserves(R.s'id = S.sid \wedge R.b'id = 103$
$\wedge Psname = 8.snarne)\}$

This query can be read as follows: "Retrieve all sailor tuples for which there exists a tuple in Reserves having the same value in the *sid* field and with *b'id* = 103." That is, for each sailor tuple, we look for a tuple in Reserves that shows that this sailor has reserved boat 103. The answer tuple *P* contains just one field, *sname*.

(Q2) Find the names of sailors who have reserved a red boat.

$\{P \mid \exists S \in Sailors\ \exists R \in Reserves(R.sid = 5.sid \wedge P.sname = S.8name$

$\wedge \exists B \in Boats(B.llid = R.bid \wedge B.color = 'red'))\}$

This query can be read as follows: "Retrieve all sailor tuples S for which there exist tuples R in Reserves and B in Boats such that $S.sid = R.sid$, $R.bid = B.bid$, and $B.coior = 'red'$." Another way to write this query, which corresponds more closely to this reading, is as follows:

$\{P \mid \exists S \in SailoTs \ \exists R \in Reserves \ \exists B \in Boats$
$(R.sid = S.sid \wedge B.bid = R.bid \wedge B.color = 'red' \wedge P.sname = S.sname)\}$

(Q7) Find the names of sailors who have reserved at least two boats.

$\{P \mid \exists S \in Sailors \ \exists R1 \in Reserves \ \exists R2 \in Reserves$
$(S.sid = R1.sid \wedge R1.sid = R2.sid \wedge R1.bid \neq R2.bid$
$\wedge P.sname = S.sname)\}$

Contrast this query with the algebra version and see how much simpler the calculus version is. In part, this difference is due to the cumbersome renaming of fields in the algebra version, but the calculus version really is simpler.

(Q9) Find the narnes of sailors who have reserved all boats.

$\{P \mid \exists S \in Sailors \ \forall B \in Boats$
$(\exists R \in Reserves(S.sid = R.sid \wedge R.bid = B.bid \wedge P.sname = S.sname))\}$

This query was expressed using the division operator in relational algebra. Note how easily it is expressed in the calculus. The calculus query directly reflects how we might express the query in English: "Find sailors S such that for all boats B there is a Reserves tuple showing that sailor S has reserved boat B."

(Q14) Find sailors who have reserved all red boats.

$\{S \mid S \in Sailor's \wedge \forall B \in Boats$
$(B.color = 'red' \Rightarrow (\exists R \in Reserves(S.sid = R.sid \wedge R.bid = B.bid)))\}$

This query can be read as follows: For each candidate (sailor), if a boat is red, the sailor must have reserved it. That is, for a candidate sailor, a boat being red must imply that the sailor has reserved it. Observe that since we can return an entire sailor tuple as the ans\ver instead of just the sailor's name, we avoided introducing a new free variable (e.g., the variable P in the previous example) to hold the answer values. On instances B1. *R2*, and *S3*, the answer contains the Sailors tuples with *sids* 22 and 31.

We can write this query without using implication, by observing that an expression of the form p $\Rightarrow$ *q* is logically equivalent to $\neg p \vee q$:

$\{S \mid S \in Sailors \wedge \forall B \in Boats$

$$(B.coioT \neq 'red' \text{ V } (\exists R \text{ E } ReSeTVeS(S.sid = R.sid \wedge R.bid = B.bid)))\}$$

This query should be read as follows: "Find sailors S such that, for all boats B, either the boat is not red or a Reserves tuple shows that sailor S has reserved boat B."

4.3.2 Domain Relational Calculus

A **domain variable** is a variable that ranges over the values in the domain of some attribute (e.g., the variable can be assigned an integer if it appears in an attribute whose domain is the set of integers). A DRC query has the form $\{(X1,X2, \ldots ,X_n) \mid P((X1,X2, \ldots ,X_n))\}$, where each Xi is either a *domain variable* or a constant and $p((Xl, X2, \ldots ,x_n))$ denotes a **DRC formula** whose only free variables are the variables among the Xi, $1 \leq i \leq n$. The result of this query is the set of all tuples $(X1, X2, \ldots , X_n)$ for which the formula evaluates to true.

A DRC formula is defined in a manner very similar to the definition of a TRC formula. The main difference is that the variables are now domain variables. Let op denote an operator in the set $\{<, >, =, \leq, \geq, \neq\}$ and let X and Y be domain variables. An **atomic formula** in DRC is one of the following:

- $(X1, X2, \ldots , X_n)$ E Rel, where Rei is a relation with n attributes; each Xi, $1 \leq i \leq n$ is either a variable or a constant
- X op Y
- X op *constant*, or *constant* op X

A **formula** is recursively defined to be one of the following, where P and q are themselves formulas and $p(X)$ denotes a formula in which the variable X appears:

- any atomic formula
- $\neg p$, $P \wedge q$, $P \vee q$, or $p \Rightarrow q$
- $\exists X(p(X))$, where X is a domain variable
- $\forall X(p(X))$, where X is a domain variable

The reader is invited to compare this definition with the definition of TRC formulas and see how closely these two definitions correspond. We will not define the semantics of DRC formulas formally; this is left as an exercise for the reader.

Examples of DRC Queries

We now illustrate DRC through several examples. The reader is invited to compare these with the TRC versions.

*(Q*11) *Find all sailors with a rating above 7.*

$$\{\langle I, N, T, A\rangle \mid (I, N, T, A) \in Sailors \wedge T > 7\}$$

This differs from the TRC version in giving each attribute a (variable) name. The condition $\langle I, N, T, A\rangle \in Sailors$ ensures that the domain variables I, N, T, and A are restricted to be fields of the *same* tuple. In comparison with the TRC query, we can say $T > 7$ instead of $S.rating > 7$, but we must specify the tuple (I, N, T, A) in the result, rather than just S.

*(Q*1) *Find the names of sailors who have reserved boat 103.*

$$\{(N) \mid \exists I, T, A(\langle I, N, T, A\rangle \in Sailors \wedge \exists Ir, Br, D((11', Br, D) \in Reserves \wedge 11' = I \wedge Br = 103))\}$$

Note that only the *sname* field is retained in the answer and that only N is a free variable. We use the notation $\exists Ir, Br, D(\ldots)$ as a shorthand for $\exists Ir(\exists Br(\exists D(\ldots)))$. Very often, all the quantified variables appear in a single relation, as in this example. An even more compact notation in this case is $\exists\langle Ir, Br, D\rangle \in Reserves$. With this notation, which we use henceforth, the query would be as follows:

$$\{(N) \mid \exists I, T, A((I, N, T, A) \in Sailors \wedge \exists\langle Ir, Br, D\rangle \in Reserves(Ir = I \wedge Br = 103))\}$$

The comparison with the corresponding TRC formula should now be straightforward. This query can also be written as follows; note the repetition of variable I and the use of the constant 103:

$$\{(N) \mid \exists I, T, A(\langle I, N, T, A\rangle \in Sailors \wedge \exists D((1,103, D) \in Reserves))\}$$

(Q2) Find the names of sailors who have Teserved a red boat.

$$\{(N) \mid \exists I, T, A((1, N, T, A) \in Sailors \wedge \exists\langle I, Br, D\rangle \in ReseTves \wedge \exists\langle Br, BN,'Ted') \in Boats)\}$$

(Q7) Find the names of sailors who have TeseTved at least two boats.

$$\{(N) \mid \exists I, T, A((1, N, T, A) \in Sailors \wedge \exists Br1, BT2, Dl, D2((1, Brl, \mathrm{DI}) \in Reserves \wedge (1, Br2, D2) \in Reserves \wedge Brl \neq Br2))\}$$

Note how the repeated use of variable *I* ensures that the same sailor has reserved both the boats in question.

(Q9) Find the names of sailors who have Teserved all boat8.

$$\{(N) \mid \exists I, T, A((I, N, T, A) \in Sailors! \wedge \forall B, BN, C(\neg(\langle B, BN, C) \in Boats) \vee (\exists \langle Ir, Br, D) \in Reserves(I = IT! \wedge BT = B))))\}$$

This query can be read as follows: "Find all values of N such that some tuple (I, N, T, A) in Sailors satisfies the following condition: For every $\langle B, BN, C)$, either this is not a tuple in Boats or there is some tuple *(IT, BT, D)* in Reserves that proves that Sailor I has reserved boat B." The $\forall$ quantifier allows the domain variables B, BN, and C to range over all values in their respective attribute domains, and the pattern '$\neg$((*B, BN,* C) E *Boats*)V' is necessary to restrict attention to those values that appear in tuples of Boats. This pattern is common in DRC formulas, and the notation $\forall\langle B, BN,$ C) E *Boats* can be used as a shortcut instead. This is similar to the notation introduced earlier for $\exists$. With this notation, the query would be written as follows:

$$\{(N) \mid \exists I, T, A((\mathrm{I}, N, T, A) \in Sailors\,! \wedge \forall\langle B, BN, \mathrm{C}) \in Boats (\exists \langle Ir, BT, D) \in ReseTves(I = IT! \wedge BT = B)))\}$$

(Q14) Find sailoTs who have TeseTved all Ted boats.

$$\{(I, N, T, A) \mid (I, N, T, A) \in SailoTs! \wedge \forall\langle B, BN, \mathrm{C}) \in Boats (\mathrm{C} = 'red' \Rightarrow \exists\langle Ir, BT, D) \in Reserves(I = IT! \wedge Br = B))\}$$

Here, we find all sailors such that, for every red boat, there is a tuple in Reserves that shows the sailor has reserved it.

4.4 EXPRESSIVE POWER OF ALGEBRA AND CALCULUS

We presented two formal query languages for the relational model. Are they equivalent in power? Can every query that can be expressed in relational algebra also be expressed in relational calculus? The answer is yes, it can. Can every query that can be expressed in relational calculus also be expressed in relational algebra? Before we answer this question, we consider a major problem with the calculus as we presented it.

Consider the query $\{S \mid \neg(S \in Sailors)\}$. This query is syntactically correct. However, it asks for all tuples S such that S is not in (the given instance of)

Sailors. The set of such S tuples is obviously infinite, in the context of infinite domains such as the set of all integers. This simple example illustrates an *unsafe* query. It is desirable to restrict relational calculus to disallow unsafe queries.

We now sketch how calculus queries are restricted to be safe. Consider a set I of relation instances, with one instance per relation that appears in the query Q. Let $Dom(Q, I)$ be the set of all constants that appear in these relation instances I or in the formulation of the query Q itself. Since we allow only finite instances I, $Dom(Q, I)$ is also finite.

For a calculus formula Q to be considered safe, at a minimum we want to ensure that, for any given I, the set of answers for Q contains only values in $Dom(Q, I)$. While this restriction is obviously required, it is not enough. Not only do we want the set of answers to be composed of constants in $Dom(Q, I)$, we wish to *compute* the set of answers by examining only tuples that contain constants in $Dom(Q, I)$! This wish leads to a subtle point associated with the use of quantifiers $\forall$ and $\exists$: Given a TRC formula of the form $\exists R(p(R))$, we want to find all values for variable R that make this formula true by checking only tuples that contain constants in $Dom(Q, I)$. Similarly, given a TRC formula of the form $\forall R(p(R))$, we want to find any values for variable R that make this formula false by checking only tuples that contain constants in $Dom(Q, I)$.

We therefore define a *safe* TRC formula Q to be a formula such that:

1. For any given I, the set of answers for Q contains only values that are in $Dom(Q, I)$.
2. For each subexpression of the form $\exists R(p(R))$ in Q, if a tuple r (assigned to variable R) makes the formula true, then r contains only constants in $Dom(Q,I)$.
3. For each subexpression of the form $\forall R(p(R))$ in Q, if a tuple r (assigned to variable R) contains a constant that is not in $Dom(Q, I)$, then r must make the formula true.

Note that this definition is not *constructive*, that is, it does not tell us how to check if a query is safe.

The query $Q = \{S \mid \neg(S \in Sailors)\}$ is unsafe by this definition. $Dom(Q,I)$ is the set of all values that appear in (an instance I of) Sailors. Consider the instance SI shown in Figure 4.1. The answer to this query obviously includes values that do not appear in $Dom(Q,SI)$.

Returning to the question of expressiveness, we can show that every query that can be expressed using a *safe* relational calculus query can also be expressed as a relational algebra query. The expressive power of relational algebra is often used as a metric of how powerful a relational database query language is. If a query language can express all the queries that we can express in relational algebra, it is said to be **relationally complete.** A practical query language is expected to be relationally complete; in addition, commercial query languages typically support features that allow us to express some queries that cannot be expressed in relational algebra.

4.5 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What is the input to a relational query? What is the result of evaluating a query? **(Section 4.1)**
- Database systems use some variant of relational algebra to represent query evaluation plans. Explain why algebra is suitable for this purpose. **(Section 4.2)**
- Describe the selection operator. What can you say about the cardinality of the input and output tables for this operator? (That is, if the input has k tuples, what can you say about the output?) Describe the projection operator. What can you say about the cardinality of the input and output tables for this operator? **(Section 4.2.1)**
- Describe the set operations of relational algebra, including union (U), intersection (n), set-difference (-), and cross-product (x). For each, what can you say about the cardinality of their input and output tables? **(Section 4.2.2)**
- Explain how the renaming operator is used. Is it required? That is, if this operator is not allowed, is there any query that can no longer be expressed in algebra? **(Section 4.2.3)**
- Define all the variations of the join operation. Why is the join operation given special attention? Cannot we express every join operation in terms of cross-product, selection, and projection? **(Section 4.2.4)**
- Define the division operation in terms of the basic relational algebra operations. Describe a typical query that calls for division. Unlike join, the division operator is not given special treatment in database systems. Explain why. **(Section 4.2.5)**

- Relational calculus is said to be a *declarative* language, in contrast to algebra, which is a *procedural* language. Explain the distinction. **(Section 4.3)**
- How does a relational calculus query 'describe' result tuples? Discuss the subset of first-order predicate logic used in tuple relational calculus, with particular attention to universal and existential quantifiers, bound and free variables, and restrictions on the query formula. **(Section 4.3.1).**
- What is the difference between tuple relational calculus and domain relational calculus? **(Section 4.3.2).**
- What is an *unsafe* calculus query? Why is it important to avoid such queries? **(Section 4.4)**
- Relational algebra and relational calculus are said to be equivalent in expressive power. Explain what this means, and how it is related to the notion of *relational completeness*. **(Section 4.4)**

EXERCISES

Exercise 4.1 Explain the statement that relational algebra operators can be *composed.* Why is the ability to compose operators important?

Exercise 4.2 Given two relations *R1* and *R2,* where *R1* contains N1 tuples, *R2* contains N2 tuples, and N2 > N1 > 0, give the minimum and maximum possible sizes (in tuples) for the resulting relation produced by each of the following relational algebra expressions. In each case, state any assumptions about the schemas for *R1* and *R2* needed to make the expression meaningful:

(1) R1 ∪ *R2,* (2) *R1* ∩ *R2,* (3) *R1* -- *R2,* (4) *R1* x *R2,* (5) *(Ta=5(R1),* (6) $\pi a(R1)$, and *(7) R1/R2*

Exercise 4.3 Consider the following schema:

Suppliers(*sid:* integer, *sname:* string, *address:* string)
Parts(pid: integer, *pname:* string, *color:* string)
Catalog(*sid:* integer, *pid:* integer, *cost:* real)

The key fields are underlined, and the domain of each field is listed after the field name. Therefore *sid* is the key for Suppliers, *pid* is the key for Parts, and *sid* and *pid* together form the key for Catalog. The Catalog relation lists the prices charged for parts by Suppliers. Write the following queries in relational algebra, tuple relational calculus, and domain relational calculus:

1. Find the *names* of suppliers who supply some red part.
2. Find the *sids* of suppliers who supply some red or green part.
3. Find the *sids* of suppliers who supply some red part or are at 221 Packer Ave.
4. Find the *sids* of suppliers who supply some rcd part and some green part.

5. Find the *sids* of suppliers who supply every part.
6. Find the *sids* of suppliers who supply every red part.
7. Find the *sids* of suppliers who supply every red or green part.
8. Find the *sids* of suppliers who supply every red part or supply every green part.
9. Find pairs of *sids* such that the supplier with the first *sid* charges more for some part than the supplier with the second *sid*.
10. Find the *pids* of parts supplied by at least two different suppliers.
11. Find the *pids* of the most expensive parts supplied by suppliers named Yosemite Sham.
12. Find the *pids* of parts supplied by every supplier at less than $200. (If any supplier either does not supply the part or charges more than $200 for it, the part is not selected.)

Exercise 4.4 Consider the Supplier-Parts-Catalog schema from the previous question. State what the following queries compute:

1. $\pi_{sname}(\pi_{sid}(\sigma_{color='red'} Parts)$!XI *(O'cost<lOoCatalog)* !XI $Suppliers)$
2. π*sname*$(\pi_{sid}((\sigma_{\text{color='red'}} Parts)$!XI $(\sigma_{cost<}$*looCatalog)* !XI $Suppliers))$
3. $(\pi_{sname}$*((O'color='red'* $Parts)$ ⋈ *(crcost<looCatalog)* !XI *Suppl'iers))* n

$$(\pi_{sname}((\sigma_{color='green'} Parts) \bowtie (\sigma_{cost<100} Catalog) \bowtie Suppliers))$$

4. *(1fsid((crcolor='red,Parts)* !XI *(crcost<10oCatalog)* ⋈ *Suppliers))* n

$$(\pi_{sid}((\sigma_{color='green'} Parts) \bowtie$$ *(crcost<lOoCatalog)* !XI *Suppliers))*

5. $\pi_{sname}((\pi_{sid,sname}((\sigma_{color='red'} Parts) \bowtie (\sigma_{cost<100} Catalog) \bowtie Suppliers))$ ∩

$$(\pi_{sid,sname}((\sigma$$*CO!07'='green'* $Parts)$!XI $(\sigma$*cost< lOoCatalog)* ⋈ *Suppliers)))*

Exercise 4.5 Consider the following relations containing airline flight information:

Flights(fino: integer, *from:* string, *to:* string,
d·istance: integer, *departs:* time, *arrives:* time)
Aircraft(*aid:* integer, *aname:* string, *cTuisingrange:* integer)
Certified(*eid:* integer, *aid:* integer)
Employees(*eid:* integer, *ename:* string, *salary:* integer)

Note that the Employees relation describes pilots and other kinds of employees as well; every pilot is certified for some aircraft (otherwise, he or she would not qualify as a pilot), and only pilots are certified to fly.

Write the following queries in relational algebra, tuple relational calculus, and domain relational calculus. Note that some of these queries may not be expressible in relational algebra (and, therefore, also not expressible in tuple and domain relational calculus)! For such queries, informally explain why they cannot be expressed. (See the exercises at the end of Chapter 5 for additional queries over the airline schenla.)

1. Finel the *eids* of pilots certified for some Boeing aircraft.
2. Find the *names* of pilots certified for some Boeing aircraft.
3. Find the *aids* of all aircraft that. can be used on non-stop flights from Bonn to Madras.

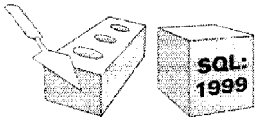
4. Identify the flights that can be piloted by every pilot whose salary is more than $100,000.
5. Find the names of pilots who can operate planes with a range greater than 3,000 miles but are not certified on any Boeing aircraft.
6. Find the *eids* of employees who make the highest salary.
7. Find the *eids* of employees who make the second highest salary.
8. Find the *eids* of employees who are certified for the largest number of aircraft.
9. Find the *eids* of employees who are certified for exactly three aircraft.
10. Find the total amount paid to employees as salaries.
11. Is there a sequence of flights from Madison to Timbuktu? Each flight in the sequence is required to depart from the city that is the destination of the previous flight; the first flight must leave Madison, the last flight must reach Timbuktu, and there is no restriction on the number of intermediate flights. Your query must determine whether a sequence of flights from Madison to Timbuktu exists for *any* input Flights relation instance.

Exercise **4.6** What is *relational completeness?* If a query language is relationally complete, can you write any desired query in that language?

Exercise **4.7** What is an *unsafe* query? Give an example and explain why it is important to disallow such queries.

BIBLIOGRAPHIC NOTES

Relational algebra was proposed by Codd in [187], and he showed the equivalence of relational algebra and TRC in [189]. Earlier, Kuhns [454] considered the use of logic to pose queries. LaCroix and Pirotte discussed DRC in [459]. Klug generalized the algebra and calculus to include aggregate operations in [439]. Extensions of the algebra and calculus to deal with aggregate functions are also discussed in [578]. Merrett proposed an extended relational algebra with quantifiers such as *the number of* that go beyond just universal and existential quantification [530]. Such generalized quantifiers are discussed at length in [52].



5

SQL: QUERIES, CONSTRAINTS, TRIGGERS

- What is included in the SQL language? What is SQL:1999?
- How are queries expressed in SQL? How is the meaning of a query specified in the SQL standard?
- How does SQL build on and extend relational algebra and calculus?
- What is grouping? How is it used with aggregate operations?
- What are nested queries?
- What are *null* values?
- How can we use queries in writing complex integrity constraints?
- What are triggers, and why are they useful? How are they related to integrity constraints?
- **Key concepts:** SQL queries, connection to relational algebra and calculus; features beyond algebra, DISTINCT clause and multiset semantics, grouping and aggregation; nested queries, correlation; set-comparison operators; *null* values, outer joins; integrity constraints specified using queries; triggers and active databases, event-condition-action rules.

What men or gods are these? What Inaiclens loth?
What mad pursuit? What struggle to escape?
\Vhat pipes and tilubrels? \Vhat wild ecstasy?

— John Keats, *Odc on a Grecian Urn*

Structured Query Language (SQL) is the most widely used conunercial relational database language. It was originally developed at IBlVI in the SEQUEL-

SQL Standards Conformance: SQL:1999 has a collection of features called Core SQL that a vendor must implement to claim conformance with the SQL:1999 standard. It is estimated that all the major vendors can comply with Core SQL with little effort. Many of the remaining features are organized into packages.

For example, packages address each of the following (with relevant chapters in parentheses): *enhanced date and time, enhanced integrity management and active databases* (this chapter), *external language interfaces* (Chapter 6), *OLAP* (Chapter 25), and *object features* (Chapter 23). The SQL/MM standard complements SQL:1999 by defining additional packages that support *data mining* (Chapter 26), *spatial data* (Chapter 28) and *text documents* (Chapter 27). Support for XML data and queries is forthcoming.

XRM and System-R projects (1974-1977). Almost immediately, other vendors introduced DBMS products based on SQL, and it is now a de facto standard. SQL continues to evolve in response to changing needs in the database area. The current ANSI/ISO standard for SQL is called SQL:1999. While not all DBMS products support the full SQL:1999 standard yet, vendors are working toward this goal and most products already support the core features. The SQL:1999 standard is very close to the previous standard, SQL-92, with respect to the features discussed in this chapter. Our presentation is consistent with both SQL-92 and SQL:1999, and we explicitly note any aspects that differ in the two versions of the standard.

5.1 OVERVIEW

The SQL language has several aspects to it.

- The Data Manipulation Language (DML): This subset of SQL allows users to pose queries and to insert, delete, and modify rows. Queries are the main focus of this chapter. We covered DML commands to insert, delete, and modify rows in Chapter 3.
- The Data Definition Language (DDL): This subset of SQL supports the creation, deletion, and modification of definitions for tables and views. *Integrity constraints* can be defined on tables, either when the table is created or later. We cocvered the DDL features of SQL in Chapter 3. Although the standard does not discuss indexes, commercial implementations also provide commands for creating and deleting indexes.
- Triggers and Advanced Integrity Constraints: The new SQL:1999 standard includes support for *triggers*, which are actions executed by the

DBMS whenever changes to the database meet conditions specified in the trigger. We cover triggers in this chapter. SQL allows the use of queries to specify complex integrity constraint specifications. We also discuss such constraints in this chapter.

- Embedded and Dynamic SQL: Embedded SQL features allow SQL code to be called from a host language such as C or COBOL. Dynamic SQL features allow a query to be constructed (and executed) at run-time. \Ve cover these features in Chapter 6.
- Client-Server Execution and Remote Database Access: These commands control how a *client* application program can connect to an SQL database *server,* or access data from a database over a network. We cover these commands in Chapter 7.
- Transaction Management: Various commands allow a user to explicitly control aspects of how a transaction is to be executed. We cover these commands in Chapter 21.
- Security: SQL provides mechanisms to control users' access to data objects such as tables and views. We cover these in Chapter 21.
- Advanced features: The SQL:1999 standard includes object-oriented features (Chapter 23), recursive queries (Chapter 24), decision support queries (Chapter 25), and also addresses emerging areas such as data mining (Chapter 26), spatial data (Chapter 28), and text and XML data management (Chapter 27).

5.1.1 Chapter Organization

The rest of this chapter is organized as follows. We present basic SQL queries in Section 5.2 and introduce SQL's set operators in Section 5.3. We discuss nested queries, in which a relation referred to in the query is itself defined within the query, in Section 5.4. We cover aggregate operators, which allow us to write SQL queries that are not expressible in relational algebra, in Section 5.5. \Ve discuss *null* values, which are special values used to indicate unknown or nonexistent field values, in Section 5.6. We discuss complex integrity constraints that can be specified using the SQL DDL in Section 5.7, extending the SQL DDL discussion from Chapter 3; the new constraint specifications allow us to fully utilize the query language capabilities of SQL.

Finally, we discuss the concept of an *active database* in Sections 5.8 and 5.9. An active database has a collection of triggers, which are specified by the DBA. A trigger describes actions to be taken when certain situations arise. The DBMS lllonitors the database, detects these situations, and invokes the trigger.

The SQL:1999 standard requires support for triggers, and several relational DBMS products already support some form of triggers.

About the Examples

We will present a number of sample queries using the following table definitions:

Sailors(*sid:* integer, *sname:* string, *rating:* integer, *age:* real)
Boats(*bid:* integer, *bname:* string, *color:* string)
Reserves(*sid:* integer, *bid:* integer, *day:* date)

We give each query a unique number, continuing with the numbering scheme used in Chapter 4. The first new query in this chapter has number Q15. Queries Q1 through Q14 were introduced in Chapter 4.[1] We illustrate queries using the instances S3 of Sailors, *R2* of Reserves, and *B1* of Boats introduced in Chapter 4, which we reproduce in Figures 5.1, 5.2, and 5.3, respectively.

All the example tables and queries that appear in this chapter are available online on the book's webpage at

http://www.cs.wisc.edu/-dbbook

The online material includes instructions on how to set up Orade, IBM DB2, Microsoft SQL Server, and MySQL, and scripts for creating the example tables and queries.

5.2 THE FORM OF A BASIC SQL QUERY

This section presents the syntax of a simple SQL query and explains its meaning through a *conceptual evaluation strategy.* A conceptual evaluation strategy is a way to evaluate the query that is intended to be easy to understand rather than efficient. A DBMS would typically execute a query in a different and more efficient way.

The basic form of an SQL query is as follows:

```
SELECT [DISTINCT] select-list
FROM   from-list
WHERE  qualification
```

[1] All references to a query can be found in the subject index for the book.

sid	sname	rating	age
22	Dustin	7	45.0
29	Brutus	1	33.0
31	Lubber	8	55.5
32	Andy	8	25.5
58	Rusty	10	35.0
64	Horatio	7	35.0
71	Zorba	10	16.0
74	Horatio	9	35.0
85	Art	3	25.5
95	Bob	3	63.5

Figure 5.1 An Instance 53 of Sailors

sid	bid	day
22	101	10/10/98
22	102	10/10/98
22	103	10/8/98
22	104	10/7/98
31	102	11/10/98
31	103	11/6/98
31	104	11/12/98
64	101	9/5/98
64	102	9/8/98
74	103	9/8/98

Figure 5.2 An Instance *R2* of Reserves

bid	bname	color
101	Interlake	blue
102	Interlake	red
103	Clipper	green
104	Marine	red

Figure 5.3 An Instance *B1* of Boats

Every query must have a SELECT clause, which specifies columns to be retained in the result, and a FROM clause, which specifies a cross-product of tables. The optional WHERE clause specifies selection conditions on the tables mentioned in the FROM clause.

Such a query intuitively corresponds to a relational algebra expression involving selections, projections, and cross-products. The close relationship between SQL and relational algebra is the basis for query optimization in a relational DBMS, as we will see in Chapters 12 and 15. Indeed, execution plans for SQL queries are represented using a variation of relational algebra expressions (Section 15.1).

Let us consider a simple example.

(Q15) Find the' names and ages of all sailors.

```
SELECT DISTINCT S.sname, S.age
FROM    Sailors S
```

The answer is a *set* of rows, each of which is a pair *(sname, age).* If two or more sailors have the same name and age, the answer still contains just one pair

with that name and age. This query is equivalent to applying the projection operator of relational algebra.

If we omit the keyword DISTINCT, we would get a copy of the row *(s,a)* for each sailor with name *s* and age *a;* the answer would be a *multiset* of rows. A **multiset** is similar to a set in that it is an unordered collection of elements, but there could be several copies of each element, and the number of copies is significant-two multisets could have the same elements and yet be different because the number of copies is different for some elements. For example, {a, b, b} and {b, a, b} denote the same multiset, and differ from the multiset {a, a, b}.

The answer to this query with and without the keyword DISTINCT on instance 53 of Sailors is shown in Figures 5.4 and 5.5. The only difference is that the tuple for Horatio appears twice if DISTINCT is omitted; this is because there are two sailors called Horatio and age 35.

snarne	*age*
Dustin	45.0
Brutus	33.0
Lubber	55.5
Andy	25.5
Rusty	35.0
Horatio	35.0
Zorba	16.0
Art	25.5
Bob	63.5

Figure 5.4 Answer to Q15

sname	*age*
Dustin	45.0
Brutus	33.0
Lubber	55.5
Andy	25.5
Rusty	35.0
Horatio	35.0
Zorba	16.0
Horatio	35.0
Art	25.5
Bob	63.5

Figure 5.5 Answer to Q15 without DISTINCT

Our next query is equivalent to an application of the selection operator of relational algebra.

*(Q*11) *Find all sailors with a rating above 7.*

```
SELECT  S.sid, S.sname, S.rating, S.age
FROM    Sailors AS S
WHERE   S.rating > 7
```

This query uses the optional keyword AS to introduce a range variable. Incidentally, when we want to retrieve all columns, as in this query, SQL provides a

convenient shorthand: We can simply write SELECT *. This notation is useful for interactive querying, but it is poor style for queries that are intended to be reused and maintained because the schema of the result is not clear from the query itself; we have to refer to the schema of the underlying Sailors table.

As these two examples illustrate, the SELECT clause is actually used to do *projection*, whereas *selections* in the relational algebra sense are expressed using the WHERE clause! This mismatch between the naming of the selection and projection operators in relational algebra and the syntax of SQL is an unfortunate historical accident.

We now consider the syntax of a basic SQL query in more detail.

- The from-list in the FROM clause is a list of table names. A table name can be followed by a range variable; a range variable is particularly useful when the same table name appears more than once in the from-list.
- The select-list is a list of (expressions involving) column names of tables named in the from-list. Column names can be prefixed by a range variable.
- The qualification in the WHERE clause is a boolean combination (i.e., an expression using the logical connectives AND, OR, and NOT) of conditions of the form *expression* op *expression*, where op is one of the comparison operators {<, <=, =, <>, >=, >}.2 An *expression* is a *column* name, a *constant*, or an (arithmetic or string) expression.
- The DISTINCT keyword is optional. It indicates that the table computed as an answer to this query should not contain *duplicates*, that is, two copies of the same row. The default is that duplicates are not eliminated.

Although the preceding rules describe (informally) the syntax of a basic SQL query, they do not tell us the *meaning* of a query. The answer to a query is itself a relation which is a *multiset* of rows in SQL!--whose contents can be understood by considering the following conceptual evaluation strategy:

1. Cmnpute the cross-product of the tables in the from-list.
2. Delete rows in the cross-product that fail the qualification conditions.
3. Delete all columns that do not appear in the select-list.
4. If DISTINCT is specified, eliminate duplicate rows.

2ExpressiollS with NOT can always be replaced by equivalent expressions without NOT given the set of comparison operators just listed.

This straightforward conceptual evaluation strategy makes explicit the rows that must be present in the answer to the query. However, it is likely to be quite inefficient. We will consider how a DB:MS actually evaluates queries in later chapters; for now, our purpose is simply to explain the meaning of a query. \Ve illustrate the conceptual evaluation strategy using the following query':

(Q1) *Find the names of sailors 'Who have reseTved boat number 103.*

It can be expressed in SQL as follows.

```
SELECT  S.sname
FROM    Sailors S, Reserves R
WHERE   S.sid = R.sid AND R.bid=103
```

Let us compute the answer to this query on the instances *R3* of Reserves and 84 of Sailors shown in Figures 5.6 and 5.7, since the computation on our usual example instances *(R2* and 83) would be unnecessarily tedious.

sid	*bid*	*day*
22	101	10/10/96
58	103	11/12/96

Figure 5.6 Instance *R3* of Reserves

sid	*sname*	*Tating*	*age*
22	dustin	7	45.0
31	lubber	8	55.5
58	rusty	10	35.0

Figure 5.7 Instance 54 of Sailors

The first step is to construct the cross-product 84 x *R3,* which is shown in Figure 5.8.

sid	*sname·j*	*rating*	*age*	*sid*	*bid*	*day*
22	dustin	7	45.0	22	101	10/10/96
22	dustin	7	45.0	58	103	11/12/96
31	lubber	8	55.5	22	101	10/10/96
31	lubber	8	55.5	58	103	11/12/96
58	rusty	10	3.5.0	22	101	10/10/96
58	rusty	10	35.0	58	103	11/12/96

Figure 5.8 *S4* x *R3*

The second step is to apply the qualification *S.sid* = *R.sid* AND *R.bid=103.* (Note that the first part of this qualification requires a join operation.) This step eliminates all but the last row from the instance shown in Figure 5.8. The third step is to eliminate unwanted columns; only *sname* appears in the SELECT clause. This step leaves us with the result shown in Figure .5.9, which is a table with a single column and, as it happens, just one row.

sname
rusty

Figure 5.9 Answer to Query Q1 011 R3 and 84

5.2.1 Examples of Basic SQL Queries

We now present several example queries, many of which were expressed earlier in relational algebra and calculus (Chapter 4). Our first example illustrates that the use of range variables is optional, unless they are needed to resolve an ambiguity. Query Q1, which we discussed in the previous section, can also be expressed as follows:

```
SELECT  sname
FROM    Sailors 5, Reserves R
WHERE   S.sid = R.sid AND bid=103
```

Only the occurrences of *sid* have to be qualified, since this column appears in both the Sailors and Reserves tables. An equivalent way to write this query is:

```
SELECT  SHame
FROM    Sailors, Reserves
WHERE   Sailors.sid = Reserves.sid AND bid=103
```

This query shows that table names can be used implicitly as row variables. Range variables need to be introduced explicitly only when the FROM clause contains more than one occurrence of a relation.[3] However, we recommend the explicit use of range variables and full qualification of all occurrences of columns with a range variable to improve the readability of your queries. We will follow this convention in all our examples.

(Q16) Find the sids of sailors who have TeseTved a red boat.

```
SELECT    R.sid
FROM      Boats B, Reserves R
WHERE     B.bid = R.bid AND 8.color = 'red'
```

This query contains a join of two tables, followed by a selection on the color of boats. We can think of 13 and R as rows in the corresponding tables that

[3]The table name cannot be used as an implicit. range variable once a range variable is introduced for t.he relation.

'prove' that a sailor with sid R.sid reserved a reel boat B.bid. On our example instances *R2* and 83 (Figures 5.1 and 5.2), the answer consists of the *sids* 22, 31, and 64. If we want the names of sailors in the result, we must also consider the Sailors relation, since Reserves does not contain this information, as the next example illustrates.

(Q2) Find the names of sailors who have reserved a Ted boat.

```
SELECT  S.sname
FROM    Sailors S, Reserves R, Boats 13
WHERE   S.sid = R.sid AND R.bid = 13.bid AND B.color = 'red'
```

This query contains a join of three tables followed by a selection on the color of boats. The join with Sailors allows us to find the name of the sailor who, according to Reserves tuple R, has reserved a red boat described by tuple 13.

(Q3) Find the coloTS of boats reseTved by LubbeT.

```
SELECT  13.color
FROM    Sailors S, Reserves R, Boats 13
WHERE   S.sid = R.sid AND R.bid = B.bid AND S.sname = 'Lubber'
```

This query is very similar to the previous one. Note that in general there may be more than one sailor called Lubber (since *sname* is not a key for Sailors); this query is still correct in that it will return the colors of boats reserved by *some* Lubber, if there are several sailors called Lubber.

(Q4) Find the names of sailors who have Teserved at least one boat.

```
SELECT  S.sname
FROM    Sailors S, Reserves R
WHERE   S.sid = R.sid
```

The join of Sailors and Reserves ensures that for each selected *sname*, the sailor has made some reservation. (If a sailor has not made a reservation, the second step in the conceptual evaluation strategy would eliminate all rows in the cross-product that involve this sailor.)

5.2.2 Expressions and Strings in the SELECT Command

SQL supports a more general version of the select-list than just a list of colu1nn8. Each item in a select-list can be of the form *expression* AS *column_name*, where *expression* is any arithmetic or string expression over column

names (possibly prefixed by range variables) and constants, and *column_name* is a new name for this column in the output of the query. It can also contain *aggregates* such as *sum* and *count*, which we will discuss in Section 5.5. The SQL standard also includes expressions over date and time values, which we will not discuss. Although not part of the SQL standard, many implementations also support the use of built-in functions such as *sqrt, sin,* and *rnod.*

(Q17) Compute increments for the mtings of peTsons who have sailed two different boats on the same day.

```
SELECT  S.sname, S.rating+1 AS rating
FROM    Sailors S, Reserves R1, Reserves R2
WHERE   S.sid = R1.sid AND S.sid = R2.sid
        AND R1.day = R2.day AND R1.bid <> R2.bid
```

Also, each item in a *qualification* can be as general as *expTession1 = expression2.*

```
SELECT  S1.sname AS name1, S2.sname AS name2
FROM    Sailors Sl, Sailors S2
WHERE   2*S1.rating = S2.rating-1
```

For string comparisons, we can use the comparison operators (=, <, >, etc.) with the ordering of strings determined alphabetically as usual. If we need to sort strings by an order other than alphabetical (e.g., sort strings denoting month names in the calendar order January, February, March, etc.), SQL supports a general concept of a collation, or sort order, for a character set. A collation allows the user to specify which characters are 'less than' which others and provides great flexibility in string manipulation.

In addition, SQL provides support for pattern matching through the LIKE operator, along with the use of the wild-card symbols % (which stands for zero or more arbitrary characters) and _ (which stands for exactly one, arbitrary, character). Thus, '_AB%' denotes a pattern matching every string that contains at least three characters, with the second and third characters being A and B respectively. Note that unlike the other comparison operators, blanks can be significant for the LIKE operator (depending on the collation for the underlying character set). Thus, *'Jeff' = 'Jeff'* is true while *'Jeff'LIKE 'Jeff '* is false. An example of the use of LIKE in a query is given below.

(Q18) Find the ages of sailors wh08e name begins and ends with B and has at least three chamcters.

```
SELECT  S.age
```

Regular Expressions in SQL: Reflecting the increased importance of text data, SQL:1999 includes a more powerful version of the LIKE operator called SIMILAR. This operator allows a rich set of regular expressions to be used as patterns while searching text. The regular expressions are similar to those supported by the Unix operating system for string searches, although the syntax is a little different.

Relational Algebra and SQL: The set operations of SQL are available in relational algebra. The main difference, of course, is that they are *multiset* operations in SQL, since tables are multisets of tuples.

```
FROM    Sailors S
WHERE   S.sname LIKE 'B.%B'
```

The only such sailor is Bob, and his age is 63.5.

5.3 UNION, INTERSECT, AND EXCEPT

SQL provides three set-manipulation constructs that extend the basic query form presented earlier. Since the answer to a query is a multiset of rows, it is natural to consider the use of operations such as union, intersection, and difference. SQL supports these operations under the names UNION, INTERSECT, and EXCEPT.[4] SQL also provides other set operations: IN (to check if an element is in a given set), op ANY, op ALL (to compare a value with the elements in a given set, using comparison operator op), and EXISTS (to check if a set is empty). IN and EXISTS can be prefixed by NOT, with the obvious modification to their meaning. We cover UNION, INTERSECT, and EXCEPT in this section, and the other operations in Section 5.4.

Consider the following query:

(Q5) Find the names of sailors who have reserved a red or a green boat.

```
SELECT  S.sname
FROM    Sailors S, Reserves R, Boats B
WHERE   S.sid = R.sid AND R.bid = B.bid
        AND (B.color = 'red' OR B.color = 'green')
```

[4]Note that although the SQL standard includes these operations, many systems currently support only UNION. Also. many systems recognize the keyword MINUS for EXCEPT.

This query is easily expressed using the OR connective in the WHERE clause. However, the following query, which is identical except for the use of 'and' rather than 'or' in the English version, turns out to be much more difficult:

(Q6) Find the names of sailors who have reserved both a red and a green boat.

If we were to just replace the use of OR in the previous query by AND, in analogy to the English statements of the two queries, we would retrieve the names of sailors who have reserved a boat that is both red and green. The integrity constraint that *bid* is a key for Boats tells us that the same boat cannot have two colors, and so the variant of the previous query with AND in place of OR will always return an empty answer set. A correct statement of Query Q6 using AND is the following:

```
SELECT S.sname
FROM   Sailors S, Reserves RI, Boats BI, Reserves R2, Boats B2
WHERE  S.sid = Rl.sid AND R1.bid = Bl.bid
       AND S.sid = R2.sid AND R2.bid = B2.bid
       AND B1.color='red' AND B2.color = 'green'
```

We can think of RI and BI as rows that prove that sailor S.sid has reserved a red boat. R2 and B2 similarly prove that the same sailor has reserved a green boat. S.sname is not included in the result unless five such rows S, RI, BI, R2, and B2 are found.

The previous query is difficult to understand (and also quite inefficient to execute, as it turns out). In particular, the similarity to the previous OR query (Query Q5) is completely lost. A better solution for these two queries is to use UNION and INTERSECT.

The OR query (Query Q5) can be rewritten as follows:

```
SELECT S.sname
FROM   Sailors S, Reserves R, Boats B
WHERE  S.sicl = R.sid AND R.bid = B.bid AND B.color = 'red'
UNION
SELECT S2.sname
FROM   Sailors S2, Boats B2, Reserves R2
WHERE  S2.sid = H2.sid AND R2.bid = B2.bicl AND B2.color = 'green'
```

This query says that we want the union of the set of sailors who have reserved red boats and the set of sailors who have reserved green boats. In complete symmetry, the AND query (Query Q6) can be rewritten as follows:

```
SELECT S.snarne
```

```
FROM    Sailors S, Reserves R, Boats B
WHERE   S.sid = R.sid AND R.bid = B.bid AND B.color = 'red'
INTERSECT
SELECT  S2.sname
FROM    Sailors S2, Boats B2, Reserves R2
WHERE   S2.sid = R2.sid AND R2.bid = B2.bid AND B2.color = 'green'
```

This query actually contains a subtle bug-if there are two sailors such as Horatio in our example instances *B1, R2,* and 83, one of whom has reserved a red boat and the other has reserved a green boat, the name Horatio is returned even though no one individual called Horatio has reserved both a red and a green boat. Thus, the query actually computes sailor names such that some sailor with this name has reserved a red boat and some sailor with the same name (perhaps a different sailor) has reserved a green boat.

As we observed in Chapter 4, the problem arises because we are using *sname* to identify sailors, and *sname* is not a key for Sailors! If we select *sid* instead of *sname* in the previous query, we would compute the set of *sids* of sailors who have reserved both red and green boats. (To compute the names of such sailors requires a nested query; we will return to this example in Section 5.4.4.)

Our next query illustrates the set-difference operation in SQL.

*(Q*19) *Find the* sids *of all sailor's who have reserved red boats but not green boats.*

```
SELECT  S.sid
FROM    Sailors S, Reserves R, Boats B
WHERE   S.sid = R.sid AND R.bid = B.bid AND B.color = 'red'
EXCEPT
SELECT  S2.sid
FROM    Sailors S2, Reserves R2, Boats B2
WHERE   S2.sid = R2.sid AND R2.bid = B2.bid AND B2.color = 'green'
```

Sailors 22, 64, and 31 have reserved red boats. Sailors 22, 74, and 31 have reserved green boats. Hence, the answer contains just the *sid 64.*

Indeed, since the Reserves relation contains sid information, there is no need to look at the Sailors relation, and we can use the following simpler query:

```
SELECT  R.sid
FROM    Boats B, Reserves R
WHERE   R.bicl = B.bid AND B.color = 'red'
EXCEPT
```

```
SELECT R2.sid
FROM   Boats B2, Reserves R2
WHERE  R2.bicl = B2.bid AND B2.color = :green'
```

Observe that this query relies on referential integrity; that is, there are no reservations for nonexisting sailors. Note that UNION, INTERSECT, and EXCEPT can be used on *any* two tables that are union-compatible, that is, have the same number of columns and the columns, taken in order, have the same types. For example, we can write the following query:

(Q20) Find all sids of sailors who have a rating of 10 or *reserved boat 104.*

```
SELECT S.sid
FROM   Sailors S
WHERE  S.rating = 10
UNION
SELECT R.sid
FROM   Reserves R
WHERE  R.bid = 104
```

The first part of the union returns the *sids* 58 and 71. The second part returns 22 and 31. The answer is, therefore, the set of *sids* 22, 31, 58, and 71. A final point to note about UNION, INTERSECT, and EXCEPT follows. In contrast to the default that duplicates are not eliminated unless DISTINCT is specified in the basic query form, the default for UNION queries is that duplicates *are* eliminated! To retain duplicates, UNION ALL must be used; if so, the number of copies of a row in the result is always m + n, where m and n are the numbers of times that the row appears in the two parts of the union. Similarly, INTERSECT ALL retains cluplicates--the number of copies of a row in the result is $min(m, n)$—and EXCEPT ALL also retains duplicates—the number of copies of a row in the result is m - n, where 'm corresponds to the first relation.

5.4 NESTED QUERIES

One of the most powerful features of SQL is nested queries. A nested query is a query that has another query embedded within it; the embedded query is called a suhquery. The embedded query can of course be a nested query itself; thus queries that have very deeply nested structures are possible. When writing a query, we sornetimes need to express a condition that refers to a table that must itself be computed. The query used to compute this subsidiary table is a subquery and appears as part of the main query. A subquery typically appears within the WHERE clause of a query. Subqueries can sometimes appear in the FROM clause or the HAVING clause (which we present in Section 5.5).

Relational Algebra and SQL: Nesting of queries is a feature that is not available in relational algebra, but nested queries can be translated into algebra, as we will see in Chapter 15. Nesting in SQL is inspired more by relational calculus than algebra. In conjunction with some of SQL's other features, such as (multi)set operators and aggregation, nesting is a very expressive construct.

This section discusses only subqueries that appear in the WHERE clause. The treatment of subqueries appearing elsewhere is quite similar. Some examples of subqueries that appear in the FROM clause are discussed later in Section 5.5.1.

5.4.1 Introduction to Nested Queries

As an example, let us rewrite the following query, which we discussed earlier, using a nested subquery:

(Q1) Find the names of sailors who have reserved boat 103.

```
SELECT S.sname
FROM   Sailors S
WHERE  S.sid IN ( SELECT R.sid
                  FROM   Reserves R
                  WHERE  R.bid = 103 )
```

The nested subquery computes the (multi)set of *sids* for sailors who have reserved boat 103 (the set contains 22,31, and 74 on instances *R2* and 83), and the top-level query retrieves the names of sailors whose *sid* is in this set. The IN operator allows us to test whether a value is in a given set of elements; an SQL query is used to generate the set to be tested. Note that it is very easy to modify this query to find all sailors who have *not* reserved boat 103-we can just replace IN by NOT IN!

The best way to understand a nested query is to think of it in terms of a conceptual evaluation strategy. In our example, the strategy consists of examining rows in Sailors and, for each such row, evaluating the subquery over Reserves. In general, the conceptual evaluation strategy that we presented for defining the semantics of a query can be extended to cover nested queries as follows: Construct the cross-product of the tables in the FROM clause of the top-level query as before. For each row in the cross-product, while testing the qllalifica-

tion in the WHERE clause, (re)compute the subquery.[5] Of course, the subquery might itself contain another nested subquery, in which case we apply the same idea one more time, leading to an evaluation strategy with several levels of nested loops.

As an example of a multiply nested query, let us rewrite the following query.

(Q2) Find the names of sailors who have reserved a red boat.

```
SELECT S.sname
FROM   Sailors S
WHERE  S.sid IN ( SELECT R.sid
                  FROM   Reserves R
                  WHERE  R.bid IN (SELECT B.bid
                                   FROM   Boats B
                                   WHERE  B.color = 'red'
```

The innermost subquery finds the set of *bids* of red boats (102 and 104 on instance *E1).* The subquery one level above finds the set of *sids* of sailors who have reserved one of these boats. On instances *E1, R2,* and 83, this set of *sids* contains 22, 31, and 64. The top-level query finds the names of sailors whose *sid* is in this set of *sids;* we get Dustin, Lubber, and Horatio.

To find the names of sailors who have not reserved a red boat, wc replace the outermost occurrence of IN by NOT IN, as illustrated in the next query.

(Q21) Find the names of sailors who have not *reserved a red boat.*

```
SELECT S.sname
FROM   Sailors S
WHERE  S.sid NOT IN ( SELECT R.sid
                      FROM   Reserves R
                      WHERE  R.bid IN ( SELECT B.bid
                                        FROM   Boats B
                                        WHERE  B.color = 'red' )
```

This qucry computes the names of sailors whose *sid* is *not* in the set 22, 31, and 64.

In contrast to Query Q21, we can modify the previous query (the nested version of Q2) by replacing the inner occurrence (rather than the outer occurence) of

[5]Since the inner subquery in our example does not depend on the 'current' row from the outer query ill any way, you rnight wonder why we have to recompute the subquery for each outer row. For an answer, see Section 5.4.2.

IN with NOT IN. This modified query would compute the names of sailors who have reserved a boat that is not red, that is, if they have a reservation, it is not for a red boat. Let us consider how. In the inner query, we check that *R.bid* is *not* either 102 or 104 (the *bids* of red boats). The outer query then finds the *sids* in Reserves tuples where the *bid* is not 102 or 104. On instances $B1$, $R2$, and 53, the outer query computes the set of *sids* 22, 31, 64, and 74. Finally, we find the names of sailors whose *sid* is in this set.

\Ve can also modify the nested query Q2 by replacing both occurrences of IN with NOT IN. This variant finds the names of sailors who have not reserved a boat that is not red, that is, who have reserved only red boats (if they've reserved any boats at all). Proceeding as in the previous paragraph, on instances *E1, R2,* and 53, the outer query computes the set of *sids* (in Sailors) other than 22, 31, 64, and 74. This is the set 29, 32, 58, 71, 85, and 95. We then find the names of sailors whose *sid* is in this set.

5.4.2 Correlated Nested Queries

In the nested queries seen thus far, the inner subquery has been completely independent of the outer query. In general, the inner subquery could depend on the row currently being examined in the outer query (in terms of our conceptual evaluation strategy). Let us rewrite the following query once more.

(Q1) Find the names of sailors who have reserved boat number 103.

```
SELECT S.sname
FROM   Sailors S
WHERE  EXISTS ( SELECT *
                FROM   Reserves R
                WHERE  R.bid = 103
                       AND R.sid = S.sid )
```

The EXISTS operator is another set comparison operator, such as IN. It allows us to test whether a set is nonempty, an implicit comparison with the empty set. Thus, for each Sailor row 5, we test whether the set of Reserves rows *R* such that *R.bid = 103* AND *S.sid = R.sid* is nonempty. If so, sailor 5 has reserved boat 103, and we retrieve the name. '1'he subquery clearly depends on the current row Sand IllUSt be re-evaluated for each row in Sailors. The occurrence of *S* in the subquery (in the form of the literal *S.sid)* is called a *cOTTelation,* and such queries are called *correlated queries.*

This query also illustrates the use of the special symbol * in situations where all we want to do is to check that a qualifying row exists, and do Hot really

want to retrieve any columns from the row. This is one of the two uses of * in the SELECT clause that is good programming style; the other is as an argument of the COUNT aggregate operation, which we describe shortly.

As a further example, by using NOT EXISTS instead of EXISTS, we can compute the names of sailors who have not reserved a red boat. Closely related to EXISTS is the UNIQUE predicate. \When we apply UNIQUE to a subquery, the resulting condition returns true if no row appears twice in the answer to the subquery, that is, there are no duplicates; in particular, it returns true if the answer is empty. (And there is also a NOT UNIQUE version.)

5.4.3 Set-Comparison Operators

We have already seen the set-comparison operators EXISTS, IN, and UNIQUE, along with their negated versions. SQL also supports op ANY and op ALL, where op is one of the arithmetic comparison operators {<, <=, =, <>, >=, >}. (SOME is also available, but it is just a synonym for ANY.)

(Q22) Find sailors whose rating is better than some sailor called Horatio.

```
SELECT  S.sid
FROM    Sailors S
WHERE   S.rating > ANY ( SELECT  S2.rating
                         FROM    Sailors S2
                         WHERE   S2.sname = 'Horatio' )
```

If there are several sailors called Horatio, this query finds all sailors whose rating is better than that of *some* sailor called Horatio. On instance 83, this computes the *sids* 31, 32, 58, 71, and 74. What if there were *no* sailor called Horatio? In this case the comparison *S.rating* > ANY ... is defined to return false, and the query returns an elnpty answer set. To understand comparisons involving ANY, it is useful to think of the comparison being carried out repeatedly. In this example, *S.rating* is successively compared with each rating value that is an answer to the nested query. Intuitively, the subquery must return a row that makes the comparison true, in order for S.*rat'ing* > ANY ... to return true.

(Q23) *Find sailors whose rating is better than every sailor' called Horat·to.*

We can obtain all such queries with a simple modification to Query Q22: Just replace ANY with ALL in the WHERE clause of the outer query. On instance $S3$, we would get the *sids* 58 and 71. If there were no sailor called Horatio, the comparison *S.rating* > ALL ... is defined to return true! The query would then return the names of all sailors. Again, it is useful to think of the comparison

being carried out repeatedly. Intuitively, the comparison must be true for every returned row for *S.rating* > ALL ... to return true.

As another illustration of ALL, consider the following query.

(Q24J Find the sailors with the highest rating.

```
SELECT  S.sid
FROM    Sailors S
WHERE   S.rating >= ALL ( SELECT  S2.rating
                          FROM    Sailors S2 )
```

The subquery computes the set of all rating values in Sailors. The outer WHERE condition is satisfied only when *S.rating* is greater than or equal to each of these rating values, that is, when it is the largest rating value. In the instance 53, the condition is satisfied only for *rating* 10, and the answer includes the *sids* of sailors with this rating, Le., 58 and 71.

Note that IN and NOT IN are equivalent to = ANY and <> ALL, respectively.

5.4.4 More Examples of Nested Queries

Let us revisit a query that we considered earlier using the INTERSECT operator.

(Q6) Find the names of sailors who have reserved both a red and a green boat.

```
SELECT  S.sname
FROM    Sailors S, Reserves R, Boats B
WHERE   S.sid = R.sid AND R.bid = B.bid AND B.color = 'red'
        AND S.sid IN ( SELECT  S2.sid
                       FROM    Sailors S2, Boats B2, Reserves R2
                       WHERE   S2.sid = R2.sid AND R2.bid = B2.bid
                               AND B2.color = 'green' )
```

This query can be understood as follows: "Find all sailors who have reserved a red boat and, further, have *sids* that are included in the set of *sids* of sailors who have reserved a green boat." This formulation of the query illustrates how queries involving INTERSECT can be rewritten using IN, which is useful to know if your system does not support INTERSECT. Queries using EXCEPT can be similarly rewritten by using NOT IN. To find the *sids* of sailors who have reserved red boats but not green boats, we can simply replace the keyword IN in the previous query by NOT IN.

As it turns out, writing this query (Q6) using INTERSECT is more complicated because we have to use *sids* to identify sailors (while intersecting) and have to return sailor names:

```
SELECT S.sname
FROM   Sailors S
WHERE  S.sid IN (( SELECT R.sid
                   FROM   Boats B, Reserves R
                   WHERE  R.bid = B.bid AND B.color = 'red' )
                   INTERSECT
                   (SELECT R2.sid
                   FROM   Boats B2, Reserves R2
                   WHERE  R2.bid = B2.bid AND B2.color = 'green' ))
```

Our next example illustrates how the *division* operation in relational algebra can be expressed in SQL.

(Q9) Find the names of sailors who have TeseTved all boats.

```
SELECT S.sname
FROM   Sailors S
WHERE  NOT EXISTS (( SELECT B.bid
                     FROM   Boats B )
                     EXCEPT
                     (SELECT R.bid
                     FROM   Reserves R
                     WHERE  R.sid = S.sid ))
```

Note that this query is correlated--for each sailor *S,* we check to see that the set of boats reserved by *S* includes every boat. An alternative way to do this query without using EXCEPT follows:

```
SELECT S.sname
FROM   Sailors S
WHERE  NOT EXISTS ( SELECT B.bid
                    FROM   Boats B
                    WHERE  NOT EXISTS ( SELECT R.bid
                                        FROM   Reserves R
                                        WHERE  R.bid = B.bid
                                               AND R.sid = S.sid ))
```

Intuitively, for each sailor we check that there is no boat that has not been reserved by this sailor.

SQL:1999 Aggregate **Functions:** The collection of aggregate functions is greatly expanded in the new standard, including several statistical functions such as standard deviation, covariance, and percentiles. However, the new aggregate functions are in the SQLjOLAP package and may not be supported by all vendors.

5.5 AGGREGATE OPERATORS

In addition to simply retrieving data, we often want to perform some computation or summarization. As we noted earlier in this chapter, SQL allows the use of arithmetic expressions. We now consider a powerful class of constructs for computing *aggregate values* such as MIN and SUM. These features represent a significant extension of relational algebra. SQL supports five aggregate operations, which can be applied on any column, say A, of a relation:

1. COUNT ([DISTINCT] A): The number of (unique) values in the A column.
2. SUM ([DISTINCT] A): The sum of all (unique) values in the A column.
3. AVG ([DISTINCT] A): The average of all (unique) values in the A column.
4. MAX (A): The maximum value in the A column.
5. MIN (A): The minimum value in the A column.

Note that it does not make sense to specify DISTINCT in conjunction with MIN or MAX (although SQL does not preclude this).

(Q25) Find the average age of all sailors.

```
SELECT AVG (S.age)
FROM   Sailors S
```

On instance 53, the average age is 37.4. Of course, the WHERE clause can be used to restrict the sailors considered in computing the average age.

(Q26) Find the average age of sailors with a rating of 10.

```
SELECT AVG (S.age)
FROM   Sailors S
WHERE  S.rating = 10
```

There are two such sailors, and their average age is 25.5. MIN (or MAX) can be used instead of AVG in the above queries to find the age of the youngest (oldest)

sailor. However) finding both the name and the age of the oldest sailor is more tricky, as the next query illustrates.

(Q27) Find the name and age of the oldest sailor.

Consider the following attempt to answer this query:

```
SELECT S.sname, MAX (S.age)
FROM   Sailors S
```

The intent is for this query to return not only the maximum age but also the name of the sailors having that age. However, this query is illegal in SQL-if the SELECT clause uses an aggregate operation, then it must use *only* aggregate operations unless the query contains a GROUP BY clause! (The intuition behind this restriction should become clear when we discuss the GROUP BY clause in Section 5.5.1.) Therefore, we cannot use MAX (S.age) as well as S.sname in the SELECT clause. We have to use a nested query to compute the desired answer to Q27:

```
SELECT S.sname, S.age
FROM   Sailors S
WHERE  S.age = ( SELECT MAX (S2.age)
                 FROM Sailors S2 )
```

Observe that we have used the result of an aggregate operation in the subquery as an argument to a comparison operation. Strictly speaking, we are comparing an age value with the result of the subquery, which is a relation. However, because of the use of the aggregate operation, the subquery is guaranteed to return a single tuple with a single field, and SQL converts such a relation to a field value for the sake of the comparison. The following equivalent query for Q27 is legal in the SQL standard but, unfortunately, is not supported in many systems:

```
SELECT S.sname, S.age
FROM   Sailors S
WHERE  ( SELECT MAX (S2.age)
         FROM Sailors S2 ) = S.age
```

We can count the number of sailors using COUNT. This example illustrates the use of * as an argument to COUNT, which is useful when \ve want to count all rows.

(Q28) Count the number of sailors.

```
SELECT COUNT (*)
```

```
FROM    Sailors S
```

We can think of * as shorthand for all the columns (in the cross-product of the **from-list** in the FROM clause). Contrast this query with the following query, which computes the number of distinct sailor names. (Remember that *sname* is not a key!)

(Q29) Count the number of different sailor names.

```
SELECT  COUNT ( DISTINCT  S.sname )
FROM    Sailors S
```

On instance 83, the answer to Q28 is 10, whereas the answer to Q29 is 9 (because two sailors have the same name, Horatio). If DISTINCT is omitted, the answer to Q29 is 10, because the name Horatio is counted twice. If COUNT does not include DISTINCT, then COUNT (*) gives the same answer as COUNT *(x)* , where *x* is any set of attributes. In our example, without DISTINCT Q29 is equivalent to Q28. However, the use of COUNT (*) is better querying style, since it is immediately clear that all records contribute to the total count.

Aggregate operations offer an alternative to the ANY and ALL constructs. For example, consider the following query:

(Q30) Find the names of sailors who are older than the oldest sailor with a rating of 10.

```
SELECT  S.sname
FROM    Sailors S
WHERE   S.age > ( SELECT  MAX  ( S2.age )
                  FROM    Sailors S2
                  WHERE   S2.rating = 10 )
```

On instance 83, the oldest sailor with rating 10 is sailor 58, whose age is 35. The names of older sailors are Bob, Dustin, Horatio, and Lubber. Using ALL, this query could alternatively be written as follows:

```
SELECT  S.sname
FROM    Sailors S
WHERE   S.age > ALL  ( SELECT  S2.age
                       FROM    Sailors S2
                       WHERE   S2.rating = 10 )
```

However, the ALL query is more error proncone could easily (and incorrectly!) use ANY instead of ALL, and retrieve sailors who are older than *some* sailor with

Relational Algebra and SQL: Aggregation is a fundamental operation that canIlot be expressed in relational algebra. Similarly, SQL's grouping construct cannot be expressed in algebra.

a rating of 10. The use of ANY intuitively corresponds to the use of MIN, instead of MAX, in the previous query.

5.5.1 The GROUP BY and HAVING Clauses

Thus far, we have applied aggregate operations to all (qualifying) rows in a relation. Often we want to apply aggregate operations to each of a number of groups of rows in a relation, where the number of groups depends on the relation instance (i.e., is not known in advance). For example, consider the following query.

(Q31) Find the age of the youngest sailor for each rating level.

If we know that ratings are integers in the range 1 to la, we could write 10 queries of the form:

```
SELECT MIN (S.age)
FROM   Sailors S
WHERE  S.rating = i
```

where i = 1,2, ... ,10. Writing 10 such queries is tedious. More important, we may not know what rating levels exist in advance.

To write such queries, we need a major extension to the basic SQL query form, namely, the GROUP BY clause. In fact, the extension also includes an optional HAVING clause that can be used to specify qualificatioIls over groups (for example, we may be interested only in rating levels> 6. The general form of an SQL query with these extensions is:

```
SELECT    [ DISTINCT] select-list
FROM      from-list
WHERE     'qualification
GROUP BY  grouping-list
HAVING    group-qualification
```

Using the GROUP BY clause, we can write Q31 as follows:

```
SELECT    S.rating, MIN (S.age)
```

```
FROM      Sailors S
GROUP BY  S.rating
```

Let us consider some important points concerning the new clauses:

- The select-list in the SELECT clause consists of (1) a list of column names and (2) a list of terms having the form aggop (*column-name*) AS *new-name*. We already saw AS used to rename output columns. Columns that are the result of aggregate operators do not already have a column name, and therefore giving the column a name with AS is especially useful.

  Every column that appears in (1) must also appear in grouping-list. The reason is that each row in the result of the query corresponds to one *group*, which is a collection of rows that agree on the values of columns in grouping-list. In general, if a column appears in list (1), but not in grouping-list, there can be multiple rows within a group that have different values in this column, and it is not clear what value should be assigned to this column in an answer row.

  We can sometimes use primary key information to verify that a column has a unique value in all rows within each group. For example, if the grouping-list contains the primary key of a table in the from-list, every column of that table has a unique value within each group. In SQL:1999, such columns are also allowed to appear in part (1) of the select-list.

- The expressions appearing in the group-qualification in the HAVING clause must have a *single* value per group. The intuition is that the HAVING clause determines whether an answer row is to be generated for a given group. To satisfy this requirement in SQL-92, a column appearing in the group-qualification must appear as the argument to an aggregation operator, or it must also appear in grouping-list. In SQL:1999, two new set functions have been introduced that allow us to check whether *every* or *any* row in a group satisfies a condition; this allows us to use conditions similar to those in a WHERE clause.

- If GROUP BY is omitted, the entire table is regarded as a single group.

We explain the semantics of such a query through an example.

(Q32) Find the age of the youngest sailor who is eligible to vote (i.e., is at least 18 years old) for each rating level with at least two such sailors.

```
SELECT    S.rating, MIN (S.age) AS minage
FROM      Sailors S
WHERE     S.age >= 18
GROUP BY  S.rating
HAVING    COUNT (*) > 1
```

We will evaluate this query on instance 83 of Sailors, reproduced in Figure 5.10 for convenience. The instance of Sailors on which this query is to be evaluated is shown in Figure 5.10. Extending the conceptual evaluation strategy presented in Section 5.2, we proceed as follows. The first step is to construct the cross-product of tables in the from-list. Because the only relation in the from-list in Query Q32 is Sailors, the result is just the instance shown in Figure 5.10.

sid	*sname*	*rating*	*age*
22	Dustin	7	45.0
29	Brutus	1	33.0
31	Lubber	8	55.5
32	Andy	8	25.5
58	Rusty	10	35.0
64	Horatio	7	35.0
71	Zorba	10	16.0
74	Horatio	9	35.0
85	Art	3	25.5
95	Bob	3	63.5
96	Frodo	3	25.5

Figure 5.10 Instance 53 of Sailors

The second step is to apply the qualification in the WHERE clause, *S.age* >= 18. This step eliminates the row (71, *zorba,* 10, 16). The third step is to eliminate unwanted columns. Only columns mentioned in the SELECT clause, the GROUP BY clause, or the HAVING clause are necessary, which means we can eliminate *sid* and *sname* in our example. The result is shown in Figure 5.11. Observe that there are two identical rows with *rating* 3 and *age* 25.5-SQL does not eliminate duplicates except when required to do so by use of the DISTINCT keyword! The number of copies of a row in the intermediate table of Figure 5.11 is determined by the number of rows in the original table that had these values in the projected columns.

The fourth step is to sort the table according to the GROUP BY clause to identify the groups. The result of this step is shown in Figure 5.12.

The fifth step is to apply the group-qualification in the HAVING clause, that is, the condition COUNT (*) > 1. This step eliminates the groups with *rating* equal to 1, 9, and 10. Observe that the order in which the WHERE and GROUP BY clauses are considered is significant: If the WHERE clause were not considered first, the group with *rating=10* would have met the group-qualification in the HAVING clause. The sixth step is to generate one answer row for each remaining group. The answer row corresponding to a group consists of a subset

rating	age
7	45.0
1	33.0
8	55.5
8	25.5
10	35.0
7	35.0
9	35.0
3	25.5
3	63.5
3	25.5

Figure 5.11 After Evaluation Step 3

rating	age
11	33.0
3	25.5
3	25.5
3	63.5
7	45.0
7	35.0
8	55.5
8	25.5
9	35.0
10	35.0

Figure 5.12 After Evaluation Step 4

of the grouping columns, plus one or more columns generated by applying an aggregation operator. In our example, each answer row has a *rating* column and a *minage* column, which is computed by applying MIN to the values in the *age* column of the corresponding group. The result of this step is shown in Figure 5.13.

rating	minage
3	25.5
7	35.0
8	25.5

Figure 5.13 Final Result in Sample Evaluation

If the query contains DISTINCT in the SELECT clause, duplicates are eliminated in an additional, and final, step.

SQL:1999 has introduced two new set functions, EVERY and ANY. To illustrate these functions, we can replace the HAVING clause in our example by

```
HAVING  COUNT (*) > 1 AND EVERY ( S.age <= 60 )
```

The fifth step of the conceptual evaluation is the one affected by the change in the HAVING clause. Consider the result of the fourth step, shown in Figure 5.12. The EVERY keyword requires that every row in a group must satisfy the attached condition to meet the group-qualification. The group for *rating* 3 does meet this criterion and is dropped; the result is shown in Figure 5.14.

SQL:1999 Extensions: Two new set functions, EVERY and ANY, have been added. When they are used in the HAVING clause, the basic intuition that the clause specifies a condition to be satisfied by each group, taken as a whole, remains unchanged. However, the condition can now involve tests on individual tuples in the group, whereas it previously relied exclusively on aggregate functions over the group of tuples.

It is worth contrasting the preceding query with the following query, in which the condition on *age* is in the WHERE clause instead of the HAVING clause:

```
SELECT   S.rating, MIN (S.age) AS minage
FROM     Sailors S
WHERE    S.age >= 18 AND S.age <= 60
GROUP BY S.rating
HAVING   COUNT (*) > 1
```

Now, the result after the third step of conceptual evaluation no longer contains the row with *age* 63.5. Nonetheless, the group for *rating* 3 satisfies the condition COUNT (*) > 1, since it still has two rows, and meets the group-qualification applied in the fifth step. The final result for this query is shown in Figure 5.15.

rating	*minage*
7	45 0
8	55.5

Figure 5.14 Final Result of EVERY Query

rating	*minage*
3	25.5
7	45.0
8	55.5

Figure 5.15 Result of Alternative Query

5.5.2 More Examples of Aggregate Queries

(Q33) For each red boat; find the number of reservations for this boat.

```
SELECT   B.bid, COUNT (*) AS reservationcount
FROM     Boats B, Reserves R
WHERE    R.bid = B.bid AND B.color = 'red'
GROUP BY B.bid
```

On instances *B1* and *R2,* the answer to this query contains the two tuples (102, 3) and (104, 2).

Observe that this version of the preceding query is illegal:

```
SELECT    B.bicl, COUNT (*) AS reservationcount
FROM      Boats B, Reserves R
WHERE     R.bid = B.bid
GROUP BY  B.bid
HAVING    B.color = 'red'
```

Even though the gToup-qualification *B.coloT = 'Ted'is* single-valued per group, since the grouping attribute *bid* is a key for Boats (and therefore determines *coloT)*, SQL disallows this query.6 Only columns that appear in the GROUP BY clause can appear in the HAVING clause, unless they appear as arguments to an aggregate operator in the HAVING clause.

(Q34) Find the avemge age of sailoTs fOT each rating level that has at least two sailoTs.

```
SELECT    S.rating, AVG (S.age) AS avgage
FROM      Sailors S
GROUP BY  S.rating
HAVING    COUNT (*) > 1
```

After identifying groups based on *mting,* we retain only groups with at least two sailors. The answer to this query on instance 83 is shown in Figure 5.16.

mting	*avgage*
3	44.5
7	40.0
8	40.5
10	25.5

Figure 5.16 Q34 Answer

mting	*avgage*
3	45.5
7	40.0
8	40.5
10	35.0

Figure 5.17 Q35 Answer

rating	*avgage*
3	45.5
7	40.0
8	40.5

Figure 5.18 Q36 Answer

The following alternative formulation of Query Q34 illustrates that the HAVING clause can have a nested subquery, just like the WHERE clause. Note that we can use *S.rating* inside the nested subquery in the HAVING clause because it has a single value for the current group of sailors:

```
SELECT    S.rating, AVG ( S.age ) AS avgage
FROM      Sailors S
GROUP BY  S.rating
HAVING    1 < ( SELECT COUNT (*)
                FROM   Sailors S2
                WHERE  S.rating = S2.rating )
```

[6]This query can be easily rewritten to be legal in SQL:1999 using EVERY in the HAVING clause.

(Q35) Find the average age of sailors who aTe of voting age (i.e., at least 18 years old) for each rating level that has at least two sailors.

```
SELECT    S.rating, AVG ( S.age ) AS avgage
FROM      Sailors S
WHERE     S. age >= 18
GROUP BY  S.rating
HAVING    1 < ( SELECT COUNT (*)
                FROM   Sailors S2
                WHERE  S.rating = S2.rating )
```

In this variant of Query Q34, we first remove tuples with *age* <= 18 and group the remaining tuples by *rating*. For each group, the subquery in the HAVING clause computes the number of tuples in Sailors (without applying the selection *age* <= 18) with the same *rating* value as the current group. If a group has less than two sailors, it is discarded. For each remaining group, we output the average age. The answer to this query on instance 53 is shown in Figure 5.17. Note that the answer is very similar to the answer for Q34, with the only difference being that for the group with rating 10, we now ignore the sailor with age 16 while computing the average.

(Q36) Find the average age oj sailors who aTe of voting age (i.e., at least 18 yeaTs old) JOT each rating level that has at least two such *sailors.*

```
SELECT    S.rating, AVG ( S.age ) AS avgage
FROM      Sailors S
WHERE     S. age> 18
GROUP BY  S.rating
HAVING    1 < ( SELECT COUNT (*)
                FROM   Sailors S2
                WHERE  S.rating = S2.rating AND S2.age >= 18 )
```

This formulation of the query reflects its similarity to Q35. The answer to Q36 on instance 53 is shown in Figure 5.18. It differs from the answer to Q35 in that there is no tuple for rating 10, since there is only one tuple with rating 10 and *age* $\geq$ 18.

Query Q36 is actually very similar to Q32, as the following simpler formulation shows:

```
SELECT    S.rating, AVG ( S.age ) AS avgage
FROM      Sailors S
WHERE     S. age> 18
GROUP BY  S.rating
```

```
HAVING    COUNT (*) > 1
```

This formulation of Q36 takes advantage of the fact that the WHERE clause is applied before grouping is done; thus, only sailors with *age* > 18 are left when grouping is done. It is instructive to consider yet another way of writing this query:

```
SELECT Temp.rating, Temp.avgage
FROM   ( SELECT    S.rating, AVG ( S.age ) AS avgage,
                   COUNT (*) AS ratingcount
         FROM      Sailors S
         WHERE     S. age> 18
         GROUP BY  S.rating) AS Temp
WHERE Temp.ratingcount > 1
```

This alternative brings out several interesting points. First, the FROM clause can also contain a nested subquery according to the SQL standard.[7] Second, the HAVING clause is not needed at all. Any query with a HAVING clause can be rewritten without one, but many queries are simpler to express with the HAVING clause. Finally, when a subquery appears in the FROM clause, using the AS keyword to give it a name is necessary (since otherwise we could not express, for instance, the condition *Temp. ratingcount* > 1).

(Q37) Find those ratings for which the average age of sailors is the m'inirnum over all ratings.

We use this query to illustrate that aggregate operations cannot be nested. One might consider writing it as follows:

```
SELECT    S.rating
FROM      Sailors S
WHERE     AVG (S.age) = ( SELECT    MIN (AVG (S2.age))
                          FROM      Sailors S2
                          GROUP BY  S2.rating )
```

A little thought shows that this query will not work even if the expression MIN (AVG (S2.age)), which is illegal, were allowed. In the nested query, Sailors is partitioned into groups by rating, and the average age is computed for each rating value. For each group, applying MIN to this average age value for the group will return the same value! A correct version of this query follows. It essentially computes a temporary table containing the average age for each rating value and then finds the rating(s) for which this average age is the minimum.

7Not all commercial database systems currently support nested queries in the FROM clause.

The Relational Model and SQL: Null values are not part of the basic relational model. Like SQL's treatment of tables as multisets of tuples, this is a departure from the basic model.

```
SELECT Temp.rating, Temp.avgage
FROM   ( SELECT   S.rating, AVG (S.age) AS avgage,
         FROM     Sailors S
         GROUP BY S.rating) AS Temp
WHERE  Temp.avgage = ( SELECT MIN (Temp.avgage) FROM Temp)
```

The answer to this query on instance 53 is $\langle 10, 25.5 \rangle$.

As an exercise, consider whether the following query computes the same answer.

```
SELECT   Temp.rating, MIN (Temp.avgage )
FROM     ( SELECT   S.rating, AVG (S.age) AS avgage,
           FROM     Sailors S
           GROUP BY S.rating) AS Temp
GROUP BY Temp.rating
```

5.6 NULL VALUES

Thus far, we have assumed that column values in a row are always known. In practice column values can be unknown. For example, when a sailor, say Dan, joins a yacht club, he may not yet have a rating assigned. Since the definition for the Sailors table has a *rating* column, what row should we insert for Dan? What is needed here is a special value that denotes *unknown*. Suppose the Sailor table definition was modified to include a *rnaiden-name* column. However, only married women who take their husband's last name have a maiden name. For women who do not take their husband's name and for men, the *maiden-name* column is *inapplicable*. Again, what value do we include in this column for the row representing Dan?

SQL provides a special column value called *null* to use in such situations. We use *null* when the column value is either *unknown* or *inapplicable*. Using our Sailor table definition, we might enter the row (98. *Dan, null*, 39) to represent Dan. The presence of *null* values complicates rnany issues, and we consider the impact of *null* values on SQL in this section.

5.6.1 Comparisons Using Null Values

Consider a comparison such as *rating* = 8. If this is applied to the row for Dan, is this condition true or false? Since Dan's rating is unknown, it is reasonable to say that this comparison should evaluate to the value unknown. In fact, this is the case for the comparisons *rating* > 8 and *rating* < 8 as well. Perhaps less obviously, if we compare two *null* values using <, >, =, and so on, the result is always unknown. For example, if we have *null* in two distinct rows of the sailor relation, any comparison returns unknown.

SQL also provides a special comparison operator IS NULL to test whether a column value is *null*; for example, we can say *rating* IS NULL, which would evaluate to true on the row representing Dan. We can also say *rating* IS NOT NULL, which would evaluate to false on the row for Dan.

5.6.2 Logical Connectives AND, OR, and NOT

Now, what about boolean expressions such as *rating* = 8 OR *age* < 40 and *mting* = 8 AND *age* < 40? Considering the row for Dan again, because *age* < 40, the first expression evaluates to true regardless of the value of *rating*, but what about the second? We can only say unknown.

But this example raises an important point—once we have *null* values, we must define the logical operators AND, OR, and NOT using a *three-valued* logic in which expressions evaluate to true, false, or unknown. We extend the usual interpretations of AND, OR, and NOT to cover the case when one of the arguments is unknown as follows. The expression NOT unknown is defined to be unknown. OR of two arguments evaluates to true if either argument evaluates to true, and to unknown if one argument evaluates to false and the other evaluates to unknown. (If both arguments are false, of course, OR evaluates to false.) AND of two arguments evaluates to false if either argument evaluates to false, and to unknown if one argument evaluates to unknown and the other evaluates to true or unknown. (If both arguments are true, AND evaluates to true.)

5.6.3 Impact on SQL Constructs

Boolean expressions arise in many contexts in SQL, and the impact of *null* values must be recognized. For example, the qualification in the WHERE clause eliminates rows (in the cross-product of tables named in the FROM clause) for which the qualification does not evaluate to true. Therefore, in the presence of *null* values, any row that evaluates to false or unknown is eliminated. Eliminating rows that evaluate to unknown has a subtle but signifieant impaet on queries, especially nested queries involving EXISTS or UNIQUE.

Another issue in the presence of *null* values is the definition of when two rows in a relation instance are regarded as *duplicates*. The SQL definition is that two rows are duplicates if corresponding columns are either equal, or both contain *null*. Contrast this definition with the fact that if we compare two *null* values using =, the result is unknown! In the context of duplicates, this comparison is implicitly treated as true, which is an anomaly.

As expected, the arithmetic operations +, -, *, and / all return *null* if one of their arguments is *null*. However, nulls can cause some unexpected behavior with aggregate operations. COUNT(*) handles *'null* values just like other values; that is, they get counted. All the other aggregate operations (COUNT, SUM, AVG, MIN, MAX, and variations using DISTINCT) simply discard *null* values—thus SUM cannot be understood as just the addition of all values in the (multi)set of values that it is applied to; a preliminary step of discarding all *null* values must also be accounted for. As a special case, if one of these operators-other than COUNT-is applied to *only* null values, the result is again *null*.

5.6.4 Outer Joins

Some interesting variants of the join operation that rely on *null* values, called outer joins, are supported in SQL. Consider the join of two tables, say Sailors $\bowtie_c$ Reserves. Tuples of Sailors that do not match some row in Reserves according to the join condition c do not appear in the result. In an outer join, on the other hanel, Sailor rows without a matching Reserves row appear exactly once in the result, with the result columns inherited from Reserves assigned *null* values.

In fact, there are several variants of the outer join idea. In a left outer join, Sailor rows without a matching Reserves row appear in the result, but not vice versa. In a right outer join, Reserves rows without a matching Sailors row appear in the result, but not vice versa. In a **full** outer join, both Sailors and Reserves rows without a match appear in the result. (Of course, rows with a match always appear in the result, for all these variants, just like the usual joins, sometimes called *inner* joins, presented in Chapter 4.)

SQL allows the desired type of join to be specified in the FROM clause. For example, the following query lists *(sid, b'id)* pairs corresponding to sailors and boats they have reserved:

```
SELECT  S.sid, R.bid
FROM    Sailors S NATURAL LEFT OUTER JOIN Reserves R
```

The NATURAL keyword specifies that the join condition is equality on all common attributes (in this example, *sid),* and the WHERE clause is not required (unless

we want to specify additional, non-join conditions). On the instances of Sailors and Reserves shown in Figure 5.6, this query computes the result shown in Figure 5.19.

sid	*bid*
22	101
31	*null*
58	103

Figure 5.19 Left Outer Join of *Sailor1* and *Reserves1*

5.6.5 Disallowing Null Values

We can disallow *null* values by specifying NOT NULL as part of the field definition; for example, *sname* CHAR(20) NOT NULL. In addition, the fields in a primary key are not allowed to take on *null* values. Thus, there is an implicit NOT NULL constraint for every field listed in a PRIMARY KEY constraint.

Our coverage of *null* values is far from complete. The interested reader should consult one of the many books devoted to SQL for a more detailed treatment of the topic.

5.7 COMPLEX INTEGRITY CONSTRAINTS IN SQL

In this section we discuss the specification of complex integrity constraints that utilize the full power of SQL queries. The features discussed in this section complement the integrity constraint features of SQL presented in Chapter 3.

5.7.1 Constraints over a Single Table

We can specify complex constraints over a single table using table constraints, which have the form CHECK *conditional-expression*. For example, to ensure that *rating* must be an integer in the range 1 to 10, we could use:

```
CREATE TABLE Sailors ( sid      INTEGER,
                       sname    CHAR(10),
                       rating   INTEGER,
                       age      REAL,
                       PRIMARY KEY (sid),
                       CHECK (rating >= 1 AND rating <= 10 ))
```

To enforce the constraint that Interlake boats cannot be reserved, we could use:

```
CREATE TABLE Reserves (sid      INTEGER,
                       bid      INTEGER,
                       day      DATE,
                       FOREIGN KEY (sid) REFERENCES Sailors
                       FOREIGN KEY (bid) REFERENCES Boats
                       CONSTRAINT noInterlakeRes
                       CHECK ( 'Interlake' <>
                               ( SELECT B.bname
                                 FROM   Boats B
                                 WHERE  B.bid = Reserves.bid )))
```

When a row is inserted into Reserves or an existing row is modified, the *conditional expression* in the CHECK constraint is evaluated. If it evaluates to false, the command is rejected.

5.7.2 Domain Constraints and Distinct Types

A user can define a new domain using the CREATE DOMAIN statement, which uses CHECK constraints.

```
CREATE DOMAIN ratingval INTEGER DEFAULT 1
                        CHECK ( VALUE >= 1 AND VALUE <= 10 )
```

INTEGER is the underlying, or *source*, type for the domain ratingval, and every ratingval value must be of this type. Values in ratingval are further restricted by using a CHECK constraint; in defining this constraint, we use the keyword VALUE to refer to a value in the domain. By using this facility, we can constrain the values that belong to a domain using the full power of SQL queries. Once a domain is defined, the name of the domain can be used to restrict column values in a table; we can use the following line in a schema declaration, for example:

```
rating   ratingval
```

The optional DEFAULT keyword is used to associate a default value with a domain. If the domain ratingval is used for a column in some relation and no value is entered for this column in an inserted tuple, the default value 1 associated with ratingval is used.

SQL's support for the concept of a domain is limited in an important respect. For example, we can define two domains called SailorId and BoatId, each

SQL:1999 Distinct Types: :Many systems, e.g., Informix UDS and IBM DB2, already support this feature. With its introduction, we expect that the support for domains will be *deprecated*, and eventually eliminated, in future versions of the SQL standard. It is really just one part of a broad set of object-oriented features in SQL:1999, which we discuss in Chapter 23.

using INTEGER as the underlying type. The intent is to force a comparison of a SailorId value with a BoatId value to always fail (since they are drawn from different domains); however, since they both have the same base type, INTEGER, the comparison will succeed in SQL. This problem is addressed through the introduction of distinct types in SqL:1999:

```
CREATE TYPE ratingtype AS INTEGER
```

This statement defines a new *distinct* type called ratingtype, with INTEGER as its source type. Values of type ratingtype can be compared with each other, but they cannot be compared with values of other types. In particular, ratingtype values are treated as being distinct from values of the source type, INTEGER—we cannot compare them to integers or combine them with integers (e.g., add an integer to a ratingtype value). If we want to define operations on the new type, for example, an *average* function, we must do so explicitly; none of the existing operations on the source type carryover. We discuss how such functions can be defined in Section 23.4.1.

5.7.3 Assertions: ICs over Several Tables

Table constraints are associated with a single table, although the conditional expression in the CHECK clause can refer to other tables. Table constraints are required to hold *only* if the a,ssociated table is nonempty. Thus, when a constraint involves two or more tables, the table constraint mechanism is sometimes cumbersome and not quite what is desired. To cover such situations, SQL supports the creation of assertions, which are constraints not associated with anyone table.

As an example, suppose that we wish to enforce the constraint that the number of boats plus the number of sailors should be less than 100. (This condition Illight be required, say, to qualify as a 'small' sailing club.) We could try the following table constraint:

```
CREATE TABLE Sailors ( sid     INTEGER,
                       sname   CHAR(10),
```

```
        rating  INTEGER,
        age     REAL,
        PRIMARY KEY (sid),
        CHECK ( rating >= 1 AND rating <= 10)
        CHECK ( ( SELECT COUNT (S.sid) FROM Sailors S )
                + ( SELECT COUNT (B.bid) FROM Boats B )
                < 100 ))
```

This solution suffers from two drawbacks. It is associated with Sailors, although it involves Boats in a completely symmetric way. More important, if the Sailors table is empty, this constraint is defined (as per the semantics of table constraints) to always hold, even if we have more than 100 rows in Boats! We could extend this constraint specification to check that Sailors is nonempty, but this approach becomes cumbersome. The best solution is to create an assertion, as follows:

```
CREATE ASSERTION smallClub
CHECK (( SELECT COUNT (S.sid) FROM Sailors S )
       + ( SELECT COUNT (B.bid) FROM Boats B)
       < 100 )
```

5.8 TRIGGERS AND ACTIVE DATABASES

A trigger is a procedure that is automatically invoked by the DBMS in response to specified changes to the database, and is typically specified by the DBA. A database that has a set of associated triggers is called an active database. A trigger description contains three parts:

- Event: A change to the database that activates the trigger.
- Condition: A query or test that is run when the trigger is activated.
- Action: A procedure that is executed when the trigger is activated and its condition is true.

A trigger can be thought of as a 'daemon' that monitors a database, and is executed when the database is modified in a way that matches the *event* specification. An insert, delete, or update statement could activate a trigger, regardless of which user or application invoked the activating statement; users may not even be aware that a trigger was executed as a side effect of their program.

A *condition* in a trigger can be a true/false statement (e.g., all employee salaries are less than $100,000) or a query. A query is interpreted as *true* if the answer

set is nonempty and *false* if the query has no answers. If the condition part evaluates to true, the action associated with the trigger is executed.

A trigger *action* can examine the answers to the query in the condition part of the trigger, refer to old and new values of tuples modified by the statement activating the trigger, execute Hew queries, and make changes to the database. In fact, an action can even execute a series of data-definition commands (e.g., create new tables, change authorizations) and transaction-oriented commands (e.g., commit) or call host-language procedures.

An important issue is when the action part of a trigger executes in relation to the statement that activated the trigger. For example, a statement that inserts records into the Students table may activate a trigger that is used to maintain statistics on how many students younger than 18 are inserted at a time by a typical insert statement. Depending on exactly what the trigger does, we may want its action to execute *before* changes are made to the Students table or *afterwards:* A trigger that initializes a variable used to count the nurnber of qualifying insertions should be executed before, and a trigger that executes once per qualifying inserted record and increments the variable should be executed after each record is inserted (because we may want to examine the values in the new record to determine the action).

5.8.1 Examples of Triggers in SQL

The examples shown in Figure 5.20, written using Oracle Server syntax for defining triggers, illustrate the basic concepts behind triggers. (The SQL:1999 syntax for these triggers is similar; we will see an example using SQL:1999 syntax shortly.) The trigger called *init_count* initializes a counter variable before every execution of an INSERT statement that adds tuples to the Students relation. The trigger called *incr_count* increments the counter for each inserted tuple that satisfies the condition *age* < 18.

One of the example triggers in Figure 5.20 executes before the aetivating statement, and the other example executes after it. A trigger can also be scheduled to execute *instead of* the activating statement; or in *deferred* fashion, at the end of the transaction containing the activating statement; or in *asynchronous* fashion, as part of a separate transaction.

The example in Figure 5.20 illustrates another point about trigger execution: A user must be able to specify whether a trigger is to be executed once per modified record or once per activating statement. If the action depends on individual changed records, for example, we have to examine the *age* field of the inserted Students record to decide whether to increment the count, the trigger-

```
CREATE TRIGGER iniLeount BEFORE INSERT ON Students      1* Event *1
    DECLARE
        count INTEGER:
    BEGIN                                               1* Action */
        count := 0:
    END

CREATE TRIGGER incLcount AFTER INSERT ON Students       1* Event *1
    WHEN (new.age < 18)  1* Condition; 'new' is just-inserted tuple *1
    FOR EACH ROW
    BEGIN        1* Action; a procedure in Oracle's PL/SQL syntax *1
        count := count + 1;
    END
```

Figure 5.20 Examples Illustrating Triggers

ing event should be defined to occur for each modified record; the FOR EACH ROW clause is used to do this. Such a trigger is called a row-level trigger. On the other hand, the *iniLcount* trigger is executed just once per INSERT statement, regardless of the number of records inserted, because we have omitted the FOR EACH ROW phrase. Such a trigger is called a statement-level trigger.

In Figure 5.20, the keyword new refers to the newly inserted tuple. If an existing tuple were modified, the keywords old and new could be used to refer to the values before and after the modification. SQL:1999 also allows the action part of a trigger to refer to the *set* of changed records, rather than just one changed record at a time. For example, it would be useful to be able to refer to the set of inserted Students records in a trigger that executes once after the INSERT statement; we could count the number of inserted records with *age* < 18 through an SQL query over this set. Such a trigger is shown in Figure 5.21 and is an aJternative to the triggers shown in Figure 5.20.

The definition in Figure 5.21 uses the syntax of SQL:1999, in order to illustrate the similarities and differences with respect to the syntax used in a typical current DBMS. The keyword clause NEW TABLE enables us to give a table name (InsertedTuples) to the set of newly inserted tuples. The FOR EACH STATEMENT clause specifies a statement-level trigger and can be omitted because it is the default. This definition does not have a WHEN clause; if such a clause is included, it follows the FOR EACH STATEMENT clause, just before the action specification.

The trigger is evaluated once for each SQL statement that inserts tuples into Students, and inserts a single tuple into a table that contains statistics on mod-

ifications to database tables. The first two fields of the tuple contain constants (identifying the modified table, Students, and the kind of modifying statement, an INSERT), and the third field is the number of inserted Students tuples with *age* < 18. (The trigger in Figure 5.20 only computes the count; an additional trigger is required to insert the appropriate tuple into the statistics table.)

```
CREATE TRIGGER seLcount AFTER INSERT ON Students        /* Event */
REFERENCING NEW TABLE AS InsertedTuples
FOR EACH STATEMENT
    INSERT                                              /* Action */
        INTO StatisticsTable(ModifiedTable, ModificationType, Count)
        SELECT 'Students', 'Insert', COUNT *
        FROM InsertedTuples I
        WHERE 1.age < 18
```

Figure 5.21 Set-Oriented Trigger

5.9 DESIGNING ACTIVE DATABASES

Triggers offer a powerful mechanism for dealing with changes to a database, but they must be used with caution. The effect of a collection of triggers can be very complex, and maintaining an active database can become very difficult. Often, a judicious use of integrity constraints can replace the use of triggers.

5.9.1 Why Triggers Can Be Hard to Understand

In an active database system, when the DBMS is about to execute a statement that modifies the database, it checks whether some trigger is activated by the statement. If so, the DBMS processes the trigger by evaluating its condition part, and then (if the condition evaluates to true) executing its action part.

If a statement activates more than one trigger, the DBMS typically processes all of them, in senne arbitrary order. An important point is that the execution of the action part of a trigger could in turn activate another trigger. In particular, the execution of the action part of a trigger could again activate the sarne trigger; such triggers are called recursive triggers. The potential for such *chain* activations and the unpredictable order in which a DBMS processes activated triggers can make it difficult to understand the effect of a collection of triggers.

5.9.2 Constraints versus Triggers

A common use of triggers is to maintain database consistency, and in such cases, we should always consider whether using an integrity constraint (e.g., a foreign key constraint) achieves the same goals. The meaning of a constraint is not defined operationally, unlike the effect of a trigger. This property makes a constraint easier to understand, and also gives the DBMS more opportunities to optimize execution. A constraint also prevents the data from being made inconsistent by *any* kind of statement, whereas a trigger is activated by a specific kind of statement (INSERT, DELETE, or UPDATE). Again, this restriction makes a constraint easier to understand.

On the other hand, triggers allow us to maintain database integrity in more flexible ways, as the following examples illustrate.

- Suppose that we have a table called Orders with fields *iternid, quantity, custornerid,* and *unitprice.* When a customer places an order, the first three field values are filled in by the user (in this example, a sales clerk). The fourth field's value can be obtained from a table called Items, but it is important to include it in the Orders table to have a complete record of the order, in case the price of the item is subsequently changed. We can define a trigger to look up this value and include it in the fourth field of a newly inserted record. In addition to reducing the number of fields that the clerk has to type in, this trigger eliminates the possibility of an entry error leading to an inconsistent price in the Orders table.
- Continuing with this example, we may want to perform some additional actions when an order is received. For example, if the purchase is being charged to a credit line issued by the company, we may want to check whether the total cost of the purchase is within the current credit limit. We can use a trigger to do the check; indeed, we can even use a CHECK constraint. Using a trigger, however, allows us to implement more sophisticated policies for dealing with purchases that exceed a credit limit. For instance, we may allow purchases that exceed the limit by no more than 10% if the customer has dealt with the company for at least a year, and add the customer to a table of candidates for credit limit increases.

5.9.3 Other Uses of Triggers

Many potential uses of triggers go beyond integrity maintenance. Triggers can alert users to unusual events (as reflected in updates to the database). For example, we may want to check whether a customer placing an order has made enough purchases in the past month to qualify for an additional discount; if so, the sales clerk must be informed so that he (or she) can tell the customer

and possibly generate additional sales! \Ve can relay this information by using a trigger that checks recent purchases and prints a message if the customer qualifies for the discount.

Triggers can generate a log of events to support auditing and security checks. For example, each time a customer places an order, we can create a record with the customer's ID and current credit limit and insert this record in a customer history table. Subsequent analysis of this table might suggest candidates for an increased credit limit (e.g., customers who have never failed to pay a bill on time and who have come within 10% of their credit limit at least three times in the last month).

As the examples in Section 5.8 illustrate, we can use triggers to gather statistics on table accesses and modifications. Some database systems even use triggers internally as the basis for managing replicas of relations (Section 22.11.1). Our list of potential uses of triggers is not exhaustive; for example, triggers have also been considered for workflow management and enforcing business rules.

5.10 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are the parts of a basic SQL query? Are the input and result tables of an SQL query sets or multisets? How can you obtain a set of tuples as the result of a query? (Section 5.2)
- What are range variables in SQL? How can you give names to output columns in a query that are defined by arithmetic or string expressions? What support does SQL offer for string pattern matching? (Section 5.2)
- What operations does SQL provide over (multi)sets of tuples, and how would you use these in writing queries? (Section 5.3)
- What are nested queries? What is *correlation* in nested queries? How would you use the operators IN, EXISTS, UNIQUE, ANY, and ALL in writing nested queries? Why are they useful? Illustrate your answer by showing how to write the *division* operator in SQL. (Section 5.4)
- What aggregate operators does SQL support? (Section 5.5)
- What is *grouping*? Is there a counterpart in relational algebra? Explain this feature, and discllss the interaction of the HAVING and WHERE clauses. Mention any restrictions that mllst be satisfied by the fields that appear in the GROUP BY clause. (Section 5.5.1)

- \What are *null* values? Are they supported in the relational model, as described in Chapter 3'1 How do they affect the meaning of queries? Can primary key fields of a table contain *null* values? (Section 5.6)
- What types of SQL constraints can be specified using the query language? Can you express primary key constraints using one of these new kinds of constraints? If so, why does SQL provide for a separate primary key constraint syntax? (Section 5.7)
- What is a *trigger*, and what are its three parts? What are the differences between row-level and statement-level triggers? (Section 5.8)
- \Vhy can triggers be hard to understand? Explain the differences between triggers and integrity constraints, and describe when you would use triggers over integrity constrains and vice versa. What are triggers used for? (Section 5.9)

EXERCISES

Online material is available for all exercises in this chapter on the book's webpage at

`http://www.cs.wisc.edu/~dbbook`

This includes scripts to create tables for each exercise for use with Oracle, IBM DB2, Microsoft SQL Server, and MySQL.

Exercise 5.1 Consider the following relations:

Student(snum: integer, *sname:* string, *major:* string, *level:* string, *age:* integer)
Class(*name:* string, *meets_at:* time, *room:* string, *fid:* integer)
Enrolled(snum: integer, *cname:* string)
Faculty(*fid:* integer, *fnarne:* string, *deptid:* integer)

The meaning of these relations is straightforward; for example, Enrolled has one record per student-class pair such that the student is enrolled in the class.

Write the following queries in SQL. No duplicates should be printed in any of the ans\vers.

1. Find the nari1es of all Juniors (level = JR) who are enrolled in a class taught by 1. Teach.
2. Find the age of the oldest student who is either a History major or enrolled in a course taught by I. Teach.
3. Find the names of all classes that either meet in room R128 or have five or more students enrolled.
4. Find the Ilames of all students who are enrolled in two classes that meet at the same time.

5. Find the names of faculty members \vho teach in every room in which some class is taught.
6. Find the names of faculty members for \vhorn the combined enrollment of the courses that they teach is less than five.
7. Print the level and the average age of students for that level, for each level.
8. Print the level and the average age of students for that level, for all levels except JR.
9. For each faculty member that has taught classes only in room *R128,* print the faculty member's name and the total number of classes she or he has taught.
10. Find the names of students enrolled in the maximum number of classes.
11. Find the names of students not enrolled in any class.
12. For each age value that appears in Students, find the level value that appears most often. For example, if there are more FR level students aged 18 than SR, JR, or SO students aged 18, you should print the pair (18, FR).

Exercise 5.2 Consider the following schema:

Suppliers(*sid:* integer, *sname:* string, *address:* string)
Parts(pid: integer, *pname:* string, *color:* string)
Catalog(*sid:* integer, *pid:* integer, *cost:* real)

The Catalog relation lists the prices charged for parts by Suppliers. Write the following queries in SQL:

1. Find the *pnames* of parts for which there is some supplier.
2. Find the *snames* of suppliers who supply every part.
3. Find the *snames* of suppliers who supply every red part.
4. Find the *pnamcs* of parts supplied by Acme Widget Suppliers and no one else.
5. Find the *sids* of suppliers who charge more for some part than the average cost of that part (averaged over all the suppliers who supply that part).
6. For each part, find the *sname* of the supplier who charges the most for that part.
7. Find the *sids* of suppliers who supply only red parts.
8. Find the *sids* of suppliers who supply a red part anel a green part.
9. Find the *sids* of suppliers who supply a red part or a green part.
10. For every supplier that only supplies green parts, print the name of the supplier and the total number of parts that she supplies.
11. For every supplier that supplies a green part and a reel part, print the name and price of the most expensive part that she supplies.

Exercise 5.3 The following relations keep track of airline flight information:

Flights(.flno: integer, *from:* string, *to:* string, *di8tance:* integer,
departs: time, *arrives:* time, *price:* integer)
Aircraft(*aid:* integer, *aname:* string, *cruisingrange:* integer)
Certified(*eid:* integer, *aid:* integer)
Employees(*eid:* integer, *ename:* string, *salary:* integer)

Note that the Employees relation describes pilots and other kinds of employees as well; every pilot is certified for some aircraft, and only pilots are certified to fly. Write each of the following queries in SQL. (*Additional queries using the same schema are listed in the exercises for Chapter 4.*)

1. Find the names of aircraft such that all pilots certified to operate them earn more than $80,000.
2. For each pilot who is certified for more than three aircraft, find the *eid* and the maximum *cruisingrange* of the aircraft for which she or he is certified.
3. Find the names of pilots whose *salary* is less than the price of the cheapest route from Los Angeles to Honolulu.
4. For all aircraft with *cruisingrange* over 1000 miles, find the name of the aircraft and the average salary of all pilots certified for this aircraft.
5. Find the names of pilots certified for some Boeing aircraft.
6. Find the *aids* of all aircraft that can be used on routes from Los Angeles to Chicago.
7. Identify the routes that can be piloted by every pilot who makes more than $100,000.
8. Print the *enames* of pilots who can operate planes with *cruisingmnge* greater than 3000 miles but are not certified on any Boeing aircraft.
9. A customer wants to travel from Madison to New York with no more than two changes of flight. List the choice of departure times from Madison if the customer wants to arrive in New York by 6 p.m.
10. Compute the difference between the average salary of a pilot and the average salary of all employees (including pilots).
11. Print the name and salary of every nonpilot whose salary is more than the average salary for pilots.
12. Print the names of employees who are certified only on aircrafts with cruising range longer than 1000 miles.
13. Print the names of employees who are certified only on aircrafts with cruising range longer than 1000 miles, but on at least two such aircrafts.
14. Print the names of employees who are certified only on aircrafts with cruising range longer than 1000 miles and who are certified on some Boeing aircraft.

Exercise 5.4 Consider the following relational schema. An employee can work in more than one department; the *pct_time* field of the Works relation shows the percentage of time that a given employee works in a given department.

Emp(eid: integer, *ename:* string, *age:* integer, *salary:* real)
Works(eid: integer, *did:* integer, *pet_time:* integer)
Dept(did.: integer, *budget:* real, *managerid:* integer)

Write the following queries in SQL:

1. Print the names and ages of each employee who works in both the Hardware department and the Software department.
2. For each department with more than 20 full-time-equivalent employees (i.e., where the part-time and full-time employees add up to at least that many full-time employees), print the *did* together with the number of employees that work in that department.

sid	sname	rating	age
18	jones	3	30.0
41	jonah	6	56.0
22	ahab	7	44.0
63	moby	*null*	15.0

Figure 5.22 An Instance of Sailors

3. Print the name of each employee whose salary exceeds the budget of all of the departments that he or she works in.
4. Find the *managerids* of managers who manage only departments with budgets greater than $1 million.
5. Find the *enames* of managers who manage the departments with the largest budgets.
6. If a manager manages more than one department, he or she *controls* the sum of all the budgets for those departments. Find the *managerids* of managers who control more than $5 million.
7. Find the *managerids* of managers who control the largest amounts.
8. Find the *enames* of managers who manage only departments with budgets larger than $1 million, but at least one department with budget less than $5 million.

Exercise 5.5 Consider the instance of the Sailors relation shown in Figure 5.22.

1. Write SQL queries to compute the average rating, using AVG; the sum of the ratings, using SUM; and the number of ratings, using COUNT.
2. If you divide the sum just computed by the count, would the result be the same as the average? How would your answer change if these steps were carried out with respect to the *age* field instead of *rating*?
3. Consider the following query: *Find the names of sailors with a higher rating than all sailors with age* < 21. The following two SQL queries attempt to obtain the answer to this question. Do they both compute the result? If not, explain why. Under what conditions would they compute the same result?

```
SELECT S.sname
FROM   Sailors S
WHERE  NOT EXISTS ( SELECT *
                    FROM   Sailors S2
                    WHERE  S2.age < 21
                           AND S.rating <= S2.rating )
```

```
SELECT *
FROM   Sailors S
WHERE  S.rating > ANY (SELECT S2.rating
                       FROM   Sailors S2
                       WHERE  S2.age < 21
```

4. Consider the instance of Sailors shown in Figure 5.22. Let us define instance S1 of Sailors to consist of the first two tuples, instance S2 to be the last two tuples, and S to be the given instance.

Show the left outer join of S with itself, with the join condition being *sid=sid.*

(b) Show the right outer join of S with itself, with the join condition being *sid=sid.*

(c) Show the full outer join of S with itself, with the join condition being *S'id=sid.*

(d) Show the left outer join of S1 with S2, with the join condition being *sid=sid.*

(e) Show the right outer join of S1 with S2, with the join condition being *sid=sid.*

(f) Show the full outer join of 81 with S2, with the join condition being *sid=sid.*

Exercise 5.6 Answer the following questions:

1. Explain the term *'impedance mismatch* in the context of embedding SQL commands in a host language such as C.
2. How can the value of a host language variable be passed to an embedded SQL command?
3. Explain the WHENEVER command's use in error and exception handling.
4. Explain the need for cursors.
5. Give an example of a situation that calls for the use of embedded SQL; that is, interactive use of SQL commands is not enough, and some host lang;uage capabilities are needed.
6. Write a C program with embedded SQL commands to address your example in the previous answer.
7. Write a C program with embedded SQL commands to find the standard deviation of sailors' ages.
8. Extend the previous program to find all sailors whose age is within one standard deviation of the average age of all sailors.
9. Explain how you would write a C program to compute the transitive closure of a graph, represented as an 8QL relation Edges(*from, to),* using embedded SQL commands. (You need not write the program, just explain the main points to be dealt with.)
10. Explain the following terms with respect to cursors: *updatability, sens,itivity,* and *scrollability.*
11. Define a cursor on the Sailors relation that is updatable, scrollable, and returns answers sorted by *age.* Which fields of Sailors can such a cursor *not* update? Why?
12. Give an example of a situation that calls for dynamic 8QL; that is, even embedded SQL is not sufficient.

Exercise 5.7 Consider the following relational schema and briefly answer the questions that follow:

Emp(*eid:* integer, *ename:* string, *age:* integer, *salary:* real)
\Vorks(*eid:* integer, *did:* integer, *pct_time:* integer)
Dept(*did:* integer, *budget:* real, *managerid:* integer)

1. Define a table constraint on Emp that will ensure that every employee makes at least $10,000.
2. Define a table constraint on Dept that will ensure that all managers have $age > 30$.
3. Define an assertion on Dept that will ensure that all managers have $age > 30$. Compare this assertion with the equivalent table constraint. Explain which is better.

4. Write SQL statements to delete all information about employees whose salaries exceed that of the manager of one or more departments that they work in. Be sure to ensure that all the relevant integrity constraints are satisfied after your updates.

Exercise 5.8 Consider the following relations:

Student(*snum:* integer, *sname:* string, *rnajor:* string, *level:* string, *age:* integer)
Class(narne: string, *meets_at:* time, *roorn:* string, *fid:* integer)
Enrolled(*snum:* integer, *cnarne:* string)
Faculty(*fid:* integer, *fnarne:* string, *deptid:* integer)

The meaning of these relations is straightforward; for example, Enrolled has one record per student-class pair such that the student is enrolled in the class.

1. Write the SQL statements required to create these relations, including appropriate versions of all primary and foreign key integrity constraints.
2. Express each of the following integrity constraints in SQL unless it is implied by the primary and foreign key constraint; if so, explain how it is implied. If the constraint cannot be expressed in SQL, say so. For each constraint, state what operations (inserts, deletes, and updates on specific relations) must be monitored to enforce the constraint.
   (a) Every class has a minimum enrollment of 5 students and a maximum enrollment of 30 students.
   (b) At least one dass meets in each room.
   (c) Every faculty member must teach at least two courses.
   (d) Only faculty in the department with *deptid=33* teach more than three courses.
   (e) Every student must be enrolled in the course called lVlathlOl.
   (f) The room in which the earliest scheduled class (i.e., the class with the smallest *meets_at* value) meets should not be the same as the room in which the latest scheduled class meets.
   (g) Two classes cannot meet in the same room at the same time.
   (h) The department with the most faculty members must have fewer than twice the number of faculty members in the department with the fewest faculty members.
   (i) No department can have more than 10 faculty members.
   (j) A student cannot add more than two courses at a time (i.e., in a single update).
   (k) The number of CS majors must be more than the number of Math majors.
   (l) The number of distinct courses in which CS majors are enrolled is greater than the number of distinct courses in which Math majors are enrolled.
   (rn) The total enrollment in courses taught by faculty in the department with *deptid=33* is greater than the number of ivlath majors.
   (n) There lllUst be at least one CS major if there are any students whatsoever.
   (0) Faculty members from different departments cannot teach in the same room.

Exercise 5.9 Discuss the strengths and weaknesses of the trigger mechanism. Contrast triggers with other integrity constraints supported by SQL.

Exercise **5.10** Consider the following relational schema. An employee can work in more than one department; the *pct_time* field of the \Works relation shows the percentage of time that a given employee works in a given department.

Emp(*eid:* integer, *ename:* string, *age:* integer, *salary:* real)
Works(*eid:* integer, *did:* integer, *pct_time:* integer)
Dept(*did:* integer, *budget:* real, *mana,gerid:* integer)

\Vrite SQL-92 integrity constraints (domain, key, foreign key, or CHECK constraints; or asser·· bons) or SQL:1999 triggers to ensure each of the following requirements, considered independently.

1. Employees must make a minimum salary of $1000.
2. Every manager must be also be an employee.
3. The total percentage of aU appointments for an employee must be under 100%.
4. A manager must always have a higher salary than any employee that he or she manages.
5. Whenever an employee is given a raise, the manager's salary must be increased to be at least as much.
6. Whenever an employee is given a raise, the manager's salary must be increased to be at least as much. Further, whenever an employee is given a raise, the department's budget must be increased to be greater than the sum of salaries of aU employees in the department.

PROJECT-BASED EXERCISE

Exercise **5.11** Identify the subset of SQL queries that are supported in Minibase.

BIBLIOGRAPHIC NOTES

The original version of SQL was developed as the query language for IBM's System R project, and its early development can be traced in [107, 151]. SQL has since become the most widely used relational query language, and its development is now subject to an international standardization process.

A very readable and comprehensive treatment of SQL-92 is presented by Melton and Simon in [524], and the central features of SQL:1999 are covered in [525]. We refer readers to these two books for an authoritative treatment of SQL. A short survey of the SQL:1999 standard is presented in [237]. Date offers an insightful critique of SQL in [202]. Although some of the problems have been addressed in SQL-92 and later revisions, others remain. A formal semantics for a large subset ofSQL queries is presented in [560]. SQL:1999 is the current International Organization for Standardization (ISO) and American National Standards Institute (ANSI) standard. Melton is the editor of the ANSI and ISO SQL:1999 standard, document ANSI/ISO/IEe 9075-:1999. The corresponding ISO document is ISO/lEe 9075-:1999. A successor, planned for 2003, builds on SQL:1999 SQL:2003 is close to ratification (as of June 20(2). Drafts of the SQL:2003 deliberations are available at the following URL:

ftp://sqlstandards.org/SC32/

[774] contains a collection of papers that cover the active database field. [794] includes a good in-depth introduction to active rules, covering smnantics, applications and design issues. [251] discusses SQL extensions for specifying integrity constraint checks through triggers. [123] also discusses a procedural mechanism, called an *alerter*, for monitoring a database. [185] is a recent paper that suggests how triggers might be incorporated into SQL extensions. Influential active database prototypes include Ariel [366], HiPAC [516J, ODE [18], Postgres [722], RDL [690], and Sentinel [36]. [147] compares various architectures for active database systems.

[32] considers conditions under which a collection of active rules has the same behavior, independent of evaluation order. Semantics of active databases is also studied in [285] and [792]. Designing and managing complex rule systems is discussed in [60, 225]. [142] discusses rule management using Chimera, a data model and language for active database systems.

PART II

APPLICATION DEVELOPMENT

6

DATABASE APPLICATION DEVELOPMENT

- How do application programs connect to a DBMS?
- How can applications manipulate data retrieved from a DBMS?
- How can applications modify data in a DBMS?
- What are cursors?
- What is JDBC and how is it used?
- What is SQLJ and how is it used?
- What are stored procedures?
- **Key concepts:** Embedded SQL, Dynamic SQL, cursors; JDBC, connections, drivers, ResultSets, java.sql, SQLJ; stored procedures, SQL/PSM

He profits most who serves best.

------Ivlotto for Rotary International

In Chapter 5, we looked at a wide range of SQL query constructs, treating SQL as an independent language in its own right. A relational DBMS supports an *interactive SQL* interface, and users can directly enter SQL commands. This simple approach is fine as long as the task at hand can be accomplished entirely with SQL cormnands. In practice, we often encounter situations in which we need the greater flexibility of a general-purpose programming language in addition to the data manipulation facilities provided by SQL. For example, we rnay want to integrate a database application with a nice graphical user interface, or we may want to integrate with other existing applications.

Applications that rely on the DBMS to manage data run as separate processes that connect to the DBlvIS to interact with it. Once a connection is established, SQL commands can be used to insert, delete, and modify data. SQL queries can be used to retrieve desired data. but we need to bridge an important difference in how a database system sees data and how an application program in a language like Java or C sees data: The result of a database query is a set (or multiset) or records, hut Java has no set or multiset data type. This mismatch is resolved through additional SQL constructs that allow applications to obtain a handle on a collection and iterate over the records one at a time.

We introduce Embedded SQL, Dynamic SQL, and cursors in Section 6.1. Embedded SQL allows us to access data using static SQL queries in application code (Section 6.1.1); with Dynamic SQL, we can create the queries at run-time (Section 6.1.3). Cursors bridge the gap between set-valued query answers and programming languages that do not support set-values (Section 6.1.2).

The emergence of Java as a popular application development language, especially for Internet applications, has made accessing a DBMS from Java code a particularly important topic. Section 6.2 covers JDBC, a prograruming interface that allows us to execute SQL queries from a Java program and use the results in the Java program. JDBC provides greater portability than Embedded SQL or Dynamic SQL, and offers the ability to connect to several DBMSs without recompiling the code. Section 6.4 covers SQLJ, which does the same for static SQL queries, but is easier to program in than Java, with JDBC.

Often, it is useful to execute application code at the database server, rather than just retrieve data and execute application logic in a separate process. Section 6.5 covers stored procedures, which enable application logic to be stored and executed at the database server. We conclude the chapter by discussing our B&N case study in Section 6.6.

While writing database applications, we must also keep in mind that typically many application programs run concurrently. The transaction concept, introduced in Chapter 1, is used to encapsulate the effects of an application on the database. An application can select certain transaction properties through SQL cormnands to control the degree to which it is exposed to the changes of other concurrently running applications. We touch on the transaction concept at many points in this chapter, and, in particular, cover transaction-related aspects of JDBC. A full discussion of transaction properties and SQL's support for transactions is deferred until Chapter 16.

Exainples that appear in this chapter are available online at

http://www.cs.wisc.edu/-dbbook

6.1 ACCESSING DATABASES FROM APPLICATIONS

In this section, we cover how SQL commands can be executed from within a program in a host language such as C or Java. The use of SQL commands within a host language program is called Embedded SQL. Details of Embedded SQL also depend on the host language. Although similar capabilities are supported for a variety of host languages, the syntax sometimes varies.

We first cover the basics of Embedded SQL with static SQL queries in Section 6.1.1. We then introduce cursors in Section 6.1.2. We discuss Dynamic SQL, which allows us to construct SQL queries at runtime (and execute them) in Section 6.1.3.

6.1.1 Embedded SQL

Conceptually, embedding SQL commands in a host language program is straightforward. SQL statements (i.e., not declarations) can be used wherever a statement in the host language is allowed (with a few restrictions). SQL statements must be clearly marked so that a preprocessor can deal with them before invoking the compiler for the host language. Also, any host language variables used to pass arguments into an SQL command must be declared in SQL. In particular, some special host language variables *must* be declared in SQL (so that, for example, any error conditions arising during SQL execution can be communicated back to the main application program in the host language).

There are, however, two complications to bear in mind. First, the data types recognized by SQL may not be recognized by the host language and vice versa. This mismatch is typically addressed by casting data values appropriately before passing them to or frorn SQL commands. (SQL, like other programming languages, provides an operator to cast values of alle type into values of another type.) The second complication has to do with SQL being set-oriented, and is addressed using cursors (see Section 6.1.2. Commands operate on and produce tables, which are sets

In our discussion of Embedded SQL, we assume that the host language is C for concreteness. because minor differcnces exist in how SQL statements are embedded in differcnt host languages.

Declaring Variables and Exceptions

SQL statements can refer to variables defined in the host program. Such host-language variables must be prefixed by a colon (:) in SQL statements and be declared between the commands EXEC SQL BEGIN DECLARE SECTION and EXEC

SQL END DECLARE SECTION. The declarations are similar to how they would look in a C program and, as usual in C. are separated by semicolons. For example. we can declare variables *c_sname*, *c_sid*, *c_rating*, and *c_age* (with the initial *c* used as a naming convention to emphasize that these are host language variables) as follows:

```
EXEC SQL BEGIN DECLARE SECTION
char c_sname[20];
long c_sid;
short c_rating;
float c_age;
EXEC SQL END DECLARE SECTION
```

The first question that arises is which SQL types correspond to the various C types, since we have just declared a collection of C variables whose values are intended to be read (and possibly set) in an SQL run-time environment when an SQL statement that refers to them is executed. The SQL-92 standard defines such a correspondence between the host language types and SQL types for a number of host languages. In our example, *c_snamc* has the type CHARACTER(20) when referred to in an SQL statement, *c_sid* has the type INTEGER, *c_rating* has the type SMALLINT, and *c_age* has the type REAL.

We also need some way for SQL to report what went wrong if an error condition arises when executing an SQL statement. The SQL-92 standard recognizes two special variables for reporting errors, SQLCODE and SQLSTATE. SQLCODE is the older of the two and is defined to return some negative value when an error condition arises, without specifying further just what error a particular negative integer denotes. SQLSTATE, introduced in the SQL-92 standard for the first time, associates predefined values with several common error conditions, thereby introducing some uniformity to how errors are reported. One of these two variables *must* be declared. The appropriate C type for SQLCODE is long and the appropriate C type for SQLSTATE is char[6J , that is, a character string five characters long. (Recall the null-terminator in C strings.) In this chapter, we assume that SQLSTATE is declared.

Embedding SQL Statements

All SQL staternents embedded within a host program must be clearly marked, with the details dependent on the host language; in C, SQL statements must be prefixed by EXEC SQL. An SQL statement can essentially appear in any place in the host language program where a host language statement can appear.

As a simple example, the following Embedded SQL statement inserts a row, whose column values are based on the values of the host language variables contained in it, into the Sailors relation:

```
EXEC SQL
INSERT INTO Sailors VALUES (:c_sname, :c_sid, :c_rating, :c_age);
```

Observe that a semicolon terminates the command, as per the convention for terminating statements in C.

The SQLSTATE variable should be checked for errors and exceptions after each Embedded SQL statement. SQL provides the WHENEVER command to simplify this tedious task:

```
EXEC SQL WHENEVER [SQLERROR | NOT FOUND] [ CONTINUE | GOTO st'mt ]
```

The intent is that the value of SQLSTATE should be checked after each Embedded SQL statement is executed. If SQLERROR is specified and the value of SQLSTATE indicates an exception, control is transferred to *stmt*, which is presumably responsible for error and exception handling. Control is also transferred to *stmt* if NOT FOUND is specified and the value of SQLSTATE is 02000, which denotes NO DATA.

6.1.2 Cursors

A major problem in embedding SQL statements in a host language like C is that an *impedance mismatch* occurs because SQL operates on *sets* of records, whereas languages like C do not cleanly support a set-of-records abstraction. The solution is to essentially provide a mechanism that allows us to retrieve rows one at a time from a relation.

This mechanism is called a cursor. We can declare a cursor on any relation or on any SQL query (because every query returns a set of rows). Once a cursor is declared, we can open it (which positions the cursor just before the first row); fetch the next row; move the cursor (to the next row, to the row after the next *n*, to the first row, or to the previous row, etc., by specifying additional parameters for the FETCH command); or close the cursor. Thus, a cursor essentially allows us to retrieve the rows in a table by positioning the cursor at a particular row and reading its contents.

Basic Cursor Definition and Usage

Cursors enable us to examine, in the host language program, a collection of rows computed by an Embedded SQL statement:

- \Ve usually need to open a cursor if the embedded statement is a SELECT (i.e.) a query). However, we can avoid opening a cursor if the answer contains a single row, as we see shortly.
- INSERT, DELETE, and UPDATE staternents typically require no cursor, although some variants of DELETE and UPDATE use a cursor.

As an example, we can find the name and age of a sailor, specified by assigning a value to the host variable *c_sid*, declared earlier, as follows:

```
EXEC SQL SELECT S.sname, S.age
         INTO    :c_sname, :c_age
         FROM    Sailors S
         WHERE   S.sid = :c_sid;
```

The INTO clause allows us to assign the columns of the single answer row to the host variables *c_sname* and *c_age*. Therefore, we do not need a cursor to embed this query in a host language program. But what about the following query, which computes the names and ages of all sailors with a rating greater than the current value of the host variable *c_minrating*?

```
SELECT S.sname, S.age
FROM   Sailors S
WHERE  S.rating > :c_minrating
```

This query returns a collection of rows, not just one row. 'When executed interactively, the answers are printed on the screen. If we embed this query in a C program by prefixing the cOlnmand with EXEC SQL, how can the answers be bound to host language variables? The INTO clause is inadequate because we must deal with several rows. The solution is to use a cursor:

```
DECLARE sinfo CURSOR FOR
SELECT S.sname, S.age
FROM   Sailors S
WHERE  S.rating > :c_minrating;
```

This code can be included in a C program, and once it is executed, the cursor *sinfo* is defined. Subsequently, we can open the cursor:

```
OPEN sinfo:
```

The value of *c_minrating* in the SQL query associated with the cursor is the value of this variable when we open the cursor. (The cursor declaration is processed at compile-time, and the OPEN command is executed at run-time.)

A cursor can be thought of as 'pointing' to a row in the collection of answers to the query associated with it. When a cursor is opened, it is positioned just before the first row. We can use the FETCH command to read the first row of cursor *sinfo* into host language variables:

```
FETCH sinfo INTO :c_sname, :c_age;
```

When the FETCH statement is executed, the cursor is positioned to point at the next row (which is the first row in the table when FETCH is executed for the first time after opening the cursor) and the column values in the row are copied into the corresponding host variables. By repeatedly executing this FETCH statement (say, in a while-loop in the C program), we can read all the rows computed by the query, one row at a time. Additional parameters to the FETCH command allow us to position a cursor in very flexible ways, but we do not discuss them.

How do we know when we have looked at all the rows associated with the cursor? By looking at the special variables SQLCODE or SQLSTATE, of course. SQLSTATE, for example, is set to the value 02000, which denotes NO DATA, to indicate that there are no more rows if the FETCH statement positions the cursor after the last row.

When we are done with a cursor, we can close it:

```
CLOSE sinfo;
```

It can be opened again if needed, and the value of : *c_minrating* in the SQL query associated with the cursor would be the value of the host variable *c_minrating* at that time.

Properties of Cursors

The general form of a cursor declaration is:

```
DECLARE cursorname [INSENSITIVE] [SCROLL] CURSOR
        [WITH HOLD]
        FOR some query
        [ ORDER BY order-item-list ]
        [ FOR READ ONLY | FOR UPDATE ]
```

A cursor can be declared to be a read-only cursor (FOR READ ONLY) or, if it is a cursor on a base relation or an updatable view, to be an updatable cursor (FOR UPDATE). If it is IIpdatable, simple variants of the UPDATE and

DELETE commands allow us to update or delete the row on which the cursor is positioned. For example, if *sinfa* is an updatable cursor and open, we can execute the following statement:

```
UPDATE  Sailors S
SET     S.rating = S.rating - 1
WHERE   CURRENT of sinfo;
```

This Embedded SQL statement modifies the *rating* value of the row currently pointed to by cursor *sinfa;* similarly, we can delete this row by executing the next statement:

```
DELETE  Sailors S
WHERE   CURRENT of sinfo;
```

A cursor is updatable by default unless it is a scrollable or insensitive cursor (see below), in which case it is read-only by default.

If the keyword SCROLL is specified, the cursor is scrollable, which means that variants of the FETCH command can be used to position the cursor in very flexible ways; otherwise, only the basic FETCH command, which retrieves the next row, is allowed.

If the keyword INSENSITIVE is specified, the cursor behaves as if it is ranging over a private copy of the collection of answer rows. Otherwise, and by default, other actions of some transaction could modify these rows, creating unpredictable behavior. For example, while we are fetching rows using the *sinfa* cursor, we might modify *rating* values in Sailor rows by concurrently executing the command:

```
UPDATE  Sailors S
SET     S.rating = S.rating -
```

Consider a Sailor row such that (1) it has not yet been fetched, and (2) its original *rating* value would have met the condition in the WHERE clause of the query associated with *sinfa,* but the new *rating* value does not. Do we fetch such a Sailor row? If INSENSITIVE is specified, the behavior is as if all answers were computed and stored when *sinfo* was opened; thus, the update command has no effect on the rows fetched by *sinfa* if it is executed after *sinfo* is opened. If INSENSITIVE is not specified, the behavior is implementation dependent in this situation.

A holdable cursor is specified using the WITH HOLD clause, and is not closed when the transaction is conunitted. The motivation for this cornes from long

transactions in which we access (and possibly change) a large number of rows of a table. If the transaction is aborted for any reason, the system potentially has to redo a lot of work when the transaction is restarted. Even if the transaction is not aborted, its locks are held for a long time and reduce the concurrency of the system. The alternative is to break the transaction into several smaller transactions, but remembering our position in the table between transactions (and other similar details) is complicated and error-prone. Allowing the application program to commit the transaction it initiated, while retaining its handle on the active table (i.e., the cursor) solves this problem: The application can commit its transaction and start a new transaction and thereby save the changes it has made thus far.

Finally, in what order do FETCH commands retrieve rows? In general this order is unspecified, but the optional ORDER BY clause can be used to specify a sort order. Note that columns mentioned in the ORDER BY clause cannot be updated through the cursor!

The order-item-list is a list of order-items; an order-item is a column name, optionally followed by one of the keywords ASC or DESC. Every column mentioned in the ORDER BY clause must also appear in the select-list of the query associated with the cursor; otherwise it is not clear what columns we should sort on. The keywords ASC or DESC that follow a column control whether the result should be sorted-with respect to that column-in ascending or descending order; the default is ASC. This clause is applied as the last step in evaluating the query.

Consider the query discussed in Section 5.5.1, and the answer shown in Figure 5.13. Suppose that a cursor is opened on this query, with the clause:

```
ORDER BY minage ASC, rating DESC
```

The answer is sorted first in ascending order by *minage,* and if several rows have the same *minage* value, these rows are sorted further in descending order by *rating.* The cursor would fetch the rows in the order shown in Figure 6.1.

rating	*minage*
8	25.5
3	25.5
7	35.0

Figure 6.1 Order in which Tuples Are Fetched

6.1.3 Dynamic SQL

Consider an application such as a spreadsheet or a graphical front-end that needs to access data from a DBMS. Such an application must accept commands from a user and, based on what the user needs, generate appropriate SQL statements to retrieve the necessary data. In such situations, we may not be able to predict in advance just what SQL statements need to be executed, even though there is (presumably) some algorithm by which the application can construct the necessary SQL statements once a user's command is issued.

SQL provides some facilities to deal with such situations; these are referred to as Dynamic SQL. We illustrate the two main commands, PREPARE and EXECUTE, through a simple example:

```
char c_sqlstring[] = {"DELETE FROM Sailors WHERE rating>5"};
EXEC SQL PREPARE readytogo FROM :c_sqlstring;
EXEC SQL EXECUTE readytogo;
```

The first statement declares the C variable *c_sqlstring* and initializes its value to the string representation of an SQL command. The second statement results in this string being parsed and compiled as an SQL command, with the resulting executable bound to the SQL variable *readytogo*. (Since *readytogo* is an SQL variable, just like a cursor name, it is not prefixed by a colon.) The third statement executes the command.

Many situations require the use of Dynamic SQL. However, note that the preparation of a Dynamic SQL command occurs at run-time and is run-time overhead. Interactive and Embedded SQL commands can be prepared once at compile-time and then re-executecl as often as desired. Consequently you should limit the use of Dynamic SQL to situations in which it is essential.

There are many more things to know about Dynamic SQL---how we can pass parameters from the host language program to the SQL statement being pre-parcel, for example--but we do not discuss it further.

6.2 AN INTRODUCTION TO JDBC

Embedded SQL enables the integration of SQL with a general-purpose programming language. As described in Section 6.1.1, a DBMS-specific preprocessor transforms the Embedded SQL statements into function calls in the host language. The details of this translation vary across DBMSs, and therefore even though the source code can be cOlnpiled to work with different DBMSs, the final executable works only with one specific DBMS.

ODBC and JDBC, short for Open DataBase Connectivity and Java DataBase Connectivity, also enable the integration of SQL with a general-purpose programming language. Both ODBC and JDBC expose database capabilities in a standardized way to the application programmer through an application programming interface (API). In contrast to Embedded SQL, ODBC and JDBC allow a single executable to access different DBMSs *without recompilation.* Thus, while Embedded SQL is DBMS-independent only at the source code level, applications using ODBC or JDBC are DBMS-independent at the source code level and at the level of the executable. In addition, using ODBC or JDBC, an application can access not just one DBMS but several different ones simultaneously.

ODBC and JDBC achieve portability at the level of the executable by introducing an extra level of indirection. All direct interaction with a specific DBMS happens through a DBMS-specific driver. A driver is a software program that translates the ODBC or JDBC calls into DBMS-specific calls. Drivers are loaded dynamically on demand since the DBMSs the application is going to access are known only at run-time. Available drivers are registered with a driver manager.

One interesting point to note is that a driver does not necessarily need to interact with a DBMS that understands SQL. It is sufficient that the driver translates the SQL commands from the application into equivalent commands that the DBMS understands. Therefore, in the remainder of this section, we refer to a data storage subsystem with which a driver interacts as a data source.

An application that interacts with a data source through ODBC or JDBC selects a data source, dynamically loads the corresponding driver, and establishes a connection with the data source. There is no limit on the number of open connections, and an application can have several open connections to different data sources. Each connection has transaction semantics; that is, changes from one connection are visible to other connections only after the connection has committed its changes. While a connection is opcn, transactions are executed by submitting SQL statements, retrieving results, processing errors, and finally committing or rolling back. The application disconnects from the data source to terminate the interaction.

In the remainder of this chapter, we concentrate on JDBC.

JDBC Drivers: The most up-to-date source of .IDBC drivers is the Sun JDBC Driver page at
http://industry.java.sun.com/products/jdbc/drivers
JDBC drivers are available for all major database sytems.

6.2.1 Architecture

The architecture of JDBC has four main components: the *application,* the *driver manager,* several data source specific *drivers,* and the corresponding *data SOUTces.*

The *application* initiates and terminates the connection with a data source. It sets transaction boundaries, submits SQL statements, and retrieves the results-----all through a well-defined interface as specified by the JDBC API. The primary goal of the *driver manager* is to load JDBC drivers and pass JDBC function calls from the application to the correct driver. The driver manager also handles JDBC initialization and information calls from the applications and can log all function calls. In addition, the driver manager performs·some rudimentary error checking. The *driver* establishes the connection with the data source. In addition to submitting requests and returning request results, the driver translates data, error formats, and error codes from a form that is specific to the data source into the JDBC standard. The *data source* processes commands from the driver and returns the results.

Depending on the relative location of the data source and the application, several architectural scenarios are possible. Drivers in JDBC are cla.ssified into four types depending on the architectural relationship between the application and the data source:

- Type I Bridges: This type of driver translates JDBC function calls into function calls of another API that is not native to the DBMS. An example is a JDBC-ODBC bridge; an application can use JDBC calls to access an ODBC compliant data source. The application loads only one driver, the bridge. Bridges have the advantage that it is easy to piggy-back the applica.tion onto an existing installation, and no new drivers have to be installed. But using bridges has several drawbacks. The increased number of layers between data source and application affects performance. In addition, the user is limited to the functionality that the ODBC driver supports.
- Type II Direct Translation to the Native API via Non-Java Driver: This type of driver translates JDBC function calls directly into method invocations of the API of one specific data source. The driver is

usually written using a combination of C++ and Java; it is dynamically linked and specific to the data source. This architecture performs significantly better than a JDBC-ODBC bridge. One disadvantage is that the database driver that implements the API needs to be installed on each computer that runs the application.

- Type III—Network Bridges: The driver talks over a network to a middleware server that translates the JDBC requests into DBMS-specific method invocations. In this case, the driver on the client site (Le., the network bridge) is not DBMS-specific. The JDBC driver loaded by the application can be quite small, as the only functionality it needs to implement is sending of SQL statements to the middleware server. The middleware server can then use a Type II JDBC driver to connect to the data source.
- Type IV-Direct Translation to the Native API via Java Driver: Instead of calling the DBMS API directly, the driver communicates with the DBMS through Java sockets. In this case, the driver on the client side is written in Java, but it is DBMS-specific. It translates JDBC calls into the native API of the database system. This solution does not require an intermediate layer, and since the implementation is all Java, its performance is usually quite good.

6.3 JDBC CLASSES AND INTERFACES

JDBC is a collection of Java classes and interfaces that enables database access from prograrl1s written in the Java language. It contains methods for connecting to a remote data source, executing SQL statements, examining sets of results from SQL statements, transaction management, and exception handling. The classes and interfaces are part of the `java.sql` package. Thus, all code fragments in the remainder of this section should include the statement `import java.sql.*` at the beginning of the code; we omit this statement in the remainder of this section. JDBC 2.0 also includes the `javax.sql` package, the JDBC Optional Package. The package `javax.sql` adds, among other things, the capability of connection pooling and the `RowSet` interface. We discuss connection pooling in Section 6.3.2, and the `ResultSet` interface in Section 6.3.4.

We now illustrate the individual steps that are required to submit a database query to a data source and to retrieve the results.

6.3.1 JDBC Driver Management

In JDBC, data source drivers are managed by the Drivermanager class, which maintains a list of all currently loaded drivers. The Drivermanager class has

methods registerDriver, deregisterDriver, and getDrivers to enable dynamic addition and deletion of drivers.

The first step in connecting to a data source is to load the corresponding JDBC driver. This is accomplished by using the Java mechanism for dynamically loading classes. The static method forName in the Class class returns the Java class as specified in the argument string and executes its static constructor. The static constructor of the dynamically loaded class loads an instance of the Driver class, and this Driver object registers itself with the DriverManager class.

The following Java example code explicitly loads a JDBC driver:

```
Class.forName("oracle/jdbc.driver.OracleDriver");
```

There are two other ways ofregistering a driver. We can include the driver with -Djdbc. drivers=oracle/jdbc. driver at the command line when we start the Java application. Alternatively, we can explicitly instantiate a driver, but this method is used only rarely, as the name of the driver has to be specified in the application code, and thus the application becomes sensitive to changes at the driver level.

After registering the driver, we connect to the data source.

6.3.2 Connections

A session with a data source is started through creation of a Connection object; A connection identifies a logical session with a data source; multiple connections within the same Java program can refer to different data sources or the same data source. Connections are specified through a **JDBC URL**, a URL that uses the jdbc protocol. Such a URL has the form

```
jdbc:<subprotocol>:<otherParameters>
```

The code example shown in Figure 6.2 establishes a connection to an Oracle database assuming that the strings userId and password are set to valid values.

In JDBC, connections can have different properties. For example, a connection can specify the granularity of transactions. If autocommit is set for a connection, then each SQL statement is considered to be its own transaction. If autocommit is off, then a series of statements that compose a transaction can be committed using the commit 0 method of the Connection class, or aborted using the rollbackO method. The Connection class has methods to set the

```
String uri = "jdbc:oracle:www.bookstore.com:3083"
Connection connection;
try {
    Connection connection =
        DriverManager.getConnection(urI,userId,password);
}
catch(SQLException excpt) {
    System.out.println(excpt.getMessageO);
    return;
}
```

Figure 6.2 Establishing a Connection with JDBC

JDBC Connections: Remember to close connections to data sources and return shared connections to the connection pool. Database systems have a limited number of resources available for connections, and orphan connections can often only be detected through time-outs-and while the database system is waiting for the connection to time-out, the resources used by the orphan connection are wasted.

autocommit mode (Connection. setAutoCommit) and to retrieve the current autocommit mode (getAutoCommit). The following methods are part of the Connection interface and permit setting and getting other properties:

- public int getTransactionIsolation() throws SQLExceptionand public void setTransactionlsolation(int 1) throws SQLException. These two functions get and set the current level of isolation for transactions handled in the current connection. All five SQL levels of isolation (see Section 16.6 for a full discussion) are possible, and argument *1* can be set as follows:
  - TRANSACTIONJNONE
  - TRANSACTIONJREAD.UNCOMMITTED
  - TRANSACTIONJREAD.COMMITTED
  - TRANSACTIONJREPEATABLEJREAD
  - TRANSACTION.BERIALIZABLE
- public boolean getReadOnlyO throws SQLException and public void setReadOnly(boolean readOnly) throws SQLException. These two functions allow the user to specify whether the transactions executecl through this connection are rcad only.

- public boolean isClosed() throws SQLException.
  Checks whether the current connection has already been closed.
- setAutoCommit and getAutoCommit.
  We already discussed these two functions.

Establishing a connection to a data source is a costly operation since it involves several steps, such as establishing a network connection to the data source, authentication, and allocation of resources such as memory. In case an application establishes many different connections from different parties (such as a Web server), connections are often **pooled** to avoid this overhead. A **connection pool** is a set of established connections to a data source. Whenever a new connection is needed, one of the connections from the pool is used, instead of creating a new connection to the data source.

Connection pooling can be handled either by specialized code in the application, or the optional javax.sql package, which provides functionality for connection pooling and allows us to set different parameters, such as the capacity of the pool, and shrinkage and growth rates. Most application servers (see Section 7.7.2) implement the javax.sql package or a proprietary variant.

6.3.3 Executing SQL Statements

We now discuss how to create and execute SQL statements using JDBC. In the JDBC code examples in this section, we assume that we have a Connection object named con. JDBC supports three different ways of executing statements: Statement, PreparedStatement, and CallableStatement. The Statement class is the base class for the other two statment classes. It allows us to query the data source with any static or dynamically generated SQL query. We cover the PreparedStatement class here and the CallableStatement class in Section 6.5, when we discuss stored procedures.

The PreparedStatement class dynamicaJly generates precompiled SQL statements that can be used several times; these SQL statements can have parameters, but their structure is fixed when the PreparedStatement object (representing the SQL statement) is created.

Consider the sample code using a PreparedStatment object shown in Figure 6.3. The SQL query specifies the query string, but uses '?' for the values of the parameters, which are set later using methods setString, setFloat, and setInt. The "1' placeholders can be used anywhere in SQL statements where they can be replaced with a value. Examples of places where they can appear include the WHERE clause (e.g., 'WHERE author=?'), or in SQL UPDATE and INSERT statements, as in Figure 6.3. The method setString is one way

```
// initial quantity is always zero
String sql = "INSERT INTO Books VALUES(?, 7, ?, ?, 0, 7)";
PreparedStatement pstmt = con.prepareStatement(sql);

// now instantiate the parameters with values
// assume that isbn, title, etc. are Java variables that
// contain the values to be inserted
pstmt.clearParameters();
pstmt.setString(l, isbn);
pstmt.setString(2, title);
pstmt.setString(3, author);
pstmt.setFloat(5, price);
pstmt.setInt(6, year);

int numRows = pstmt.executeUpdate();
```

Figure 6.3 SQL Update Using a PreparedStatement Object

to set a parameter value; analogous methods are available for int, float, and date. It is good style to always use clearParameters 0 before setting parameter values in order to remove any old data.

There are different ways of submitting the query string to the data source. In the example, we used the executeUpdate command, which is used if we know that the SQL statement does not return any records (SQL UPDATE, INSERT, ALTER, and DELETE statements). The executeUpdate method returns an integer indicating the number of rows the SQL statement modified; it returns 0 for successful execution without modifying any rows.

The executeQuery method is used if the SQL statement returns data, such as in a regular SELECT query. JDBC has its own cursor mechanism in the form of a ResultSet object, which we discuss next. The execute method is more general than executeQuery and executeUpdate; the references at the end of the chapter provide pointers with more details.

6.3.4 ResultSets

As discussed in the previous section, the statement executeQuery returns a ResultSet object, which is similar to a cursor. ResultSet cursors in JDBC 2.0 are very powerful; they allow forward and reverse scrolling and in-place editing and insertions.

In its most basic form, the ResultSet object allows us to read one row of the output of the query at a time. Initially, the ResultSet is positioned before the first row, and we have to retrieve the first row with an explicit call to the next() method. The next method returns false if there are no more rows in the query answer, and true other\vise. The code fragment shown in Figure 6.4 illustrates the basic usage of a ResultSet object.

```
ResultSet rs=stmt.executeQuery(sqlQuery);
// rs is now a cursor
// first call to rs.nextO moves to the first record
// rs.nextO moves to the next row
String sqlQuery;
ResultSet rs = stmt.executeQuery(sqlQuery)
while (rs.next()) {
    // process the data
}
```

Figure 6.4 Using a ResultSet Object

While next() allows us to retrieve the logically next row in the query answer, we can move about in the query answer in other ways too:

- previous() moves back one row.
- absolute(int num) moves to the row with the specified number.
- relative(int num) moves forward or backward (if num is negative) relative to the current position. relative(-1) has the same effect as previous.
- first() moves to the first row, and last() moves to the last row.

Matching Java and SQL Data Types

In considering the interaction of an application with a data source, the issues we encountered in the context of Embedded SQL (e.g., passing information between the application and the data source through shared variables) arise again. To deal with such issues, JDBC provides special data types and specifies their relationship to corresponding SQL data types. Figure 6.5 shows the accessor methods in a ResultSet object for the most common SQL datatypes. With these accessor methods, we can retrieve values from the current row of the query result referenced by the ResultSet object. There are two forms for each accessor method: One method retrieves values by column index, starting at one, and the other retrieves values by column name. The following example shows how to access fields of the current ResultSet row using accesssor methods.

SQL Type	Java class	ResultSet get method
BIT	Boolean	getBooleanO
CHAR	String	getStringO
VARCHAR	String	getStringO
DOUBLE	Double	getDoubleO
FLOAT	Double	getDoubleO
INTEGER	Integer	getIntO
REAL	Double	getFloatO
DATE	java.sql.Date	getDateO
TIME	java.sql.Time	getTimeO
TIMESTAMP	java.sql.TimeStamp	getTimestamp()

Figure 6.5 Reading SQL Datatypes from a ResultSet Object

```
ResultSet rs=stmt.executeQuery(sqIQuery);
String sqlQuerYi
ResultSet rs = stmt.executeQuery(sqIQuery)
while (rs.nextO) {
    isbn = rs.getString(l);
    title = rs.getString("TITLE");
    // process isbn and title
}
```

6.3.5 Exceptions and Warnings

Similar to the SQLSTATE variable, most of the methods in java. sql can throw an exception of the type SQLException if an error occurs. The information includes SQLState, a string that describes the error (e.g., whether the statement contained an SQL syntax error). In addition to the standard getMessage 0 method inherited from Throwable, SQLException has two additional methods that provide further information, and a method to get (or chain) additional exceptions:

- public String getSQLState 0 returns an SQLState identifier based on the SQL:1999 specification, as discussed in Section 6.1.1.
- public int getErrorCode () retrieves a vendor-specific error code.
- public SQLException getNextExceptionO gets the next exception in a chain of exceptions associated with the current SQLException object.

An SQL\¥arning is a subclass of SQLException. Warnings are not as severe as errors and the program can usually proceed without special handling of warnings. \Varnings are not thrown like other exceptions, and they are not caught as

part of the try"-catch block around a java.sql statement. We Heed to specifically test whether warnings exist. Connection, Statement, and ResultSet objects all have a getWarnings 0 method with which we can retrieve SQL warnings if they exist. Duplicate retrieval of warnings can be avoided through clearWarnings O. Statement objects clear warnings automatically on execution of the next statement; ResultSet objects clear warnings every time a new tuple is accessed.

Typical code for obtaining SQLWarnings looks similar to the code shown in Figure 6.6.

```
try {
    stmt = con.createStatement();
    warning = con.getWarnings();
    while( warning != null) {
        // handleSQLWarnings          //code to process warning
        warning = warning.getNextWarningO;      //get next warning
    }
    con.clear\Varnings();

    stmt.executeUpdate( queryString );
    warning = stmt.getWarnings();
    while( warning != null) {
        // handleSQLWarnings          //code to process warning
        warning = warning.getNextWarningO;      //get next warning
    }
} // end try
catch ( SQLException SQLe) {
    // code to handle exception
} // end catch
```

Figure 6.6 Processing JDBC Warnings and Exceptions

6.3.6 Examining Database Metadata

We can use the DatabaseMetaData object to obtain information about the database system itself, as well as information frorn the database catalog. For example, the following code fragment shows how to obtain the name and driver version of the JDBC driver:

```
DatabaseMetaData md = con.getMetaD<Lta():

System.out.println("Driver Information:");
```

```
System.out.println("Name:" + md.getDriverNameO
            + "; version:" + mcl.getDriverVersion());
```

The DatabaseMetaData object has many more methods (in JDBC 2.0, exactly 134); we list some methods here:

- public ResultSet getCatalogs0 throws SqLException. This function returns a ResultSet that can be used to iterate over all the catalog relations. The functions getIndexInfo0 and getTables0 work analogously.
- pUblic int getMaxConnections0 throws SqLException. This function returns the ma.ximum number of connections possible.

We will conclude our discussion of JDBC with an example code fragment that examines all database metadata shown in Figure 6.7.

```
DatabaseMetaData dmd = con.getMetaDataO;
ResultSet tablesRS = dmd.getTables(null,null,null,null);
string tableName;

while(tablesRS.next()) {
    tableNarne = tablesRS.getString("TABLE_NAME");

    // print out the attributes of this table
    System.out.println("The attributes of table"
            + tableName + " are:");
    ResultSet columnsRS = dmd.getColums(null,null,tableName, null);
    while (columnsRS.next()) {
        System.out.print(columnsRS.getString("COLUMN_NAME")
                +" ");
    }

    // print out the primary keys of this table
    System.out.println("The keys of table" + tableName + " are:");
    ResultSet keysRS = dmd.getPrimaryKeys(null,null,tableName);
    while (keysRS.next()) {
        'System.out.print(keysRS.getStringC'COLUMN_NAME") +" ");
    }
}
```

Figure 6.7 Obtaining Infon-nation about a Data Source

6.4 SQLJ

SQLJ (short for 'SQL-Java') was developed by the SQLJ Group, a group of database vendors and Sun. SQLJ was developed to complement the dynamic way of creating queries in JDBC with a static model. It is therefore very close to Embedded SQL. Unlike JDBC, having semi-static SQL queries allows the compiler to perform SQL syntax checks, strong type checks of the compatibility of the host variables with the respective SQL attributes, and consistency of the query with the database schema-tables, attributes, views, and stored procedures—all at compilation time. For example, in both SQLJ and Embedded SQL, variables in the host language always are bound statically to the same arguments, whereas in JDBC, we need separate statements to bind each variable to an argument and to retrieve the result. For example, the following SQLJ statement binds host language variables title, price, and author to the return values of the cursor books.

```
#sql books = {
    SELECT title, price INTO :title, :price
    FROM Books WHERE author = :author
};
```

In JDBC, we can dynamically decide which host language variables will hold the query result. In the following example, we read the title of the book into variable ftitle if the book was written by Feynman, and into variable otitle otherwise:

```
// assume we have a ResultSet cursor rs
author = rs.getString(3);

if (author=="Feynman") {
    ftitle = rs.getString(2):
}
else {
    otitle = rs.getString(2);
}
```

When writing SQLJ applications, we just write regular Java code and embed SQL statements according to a set of rules. SQLJ applications are pre-processed through an SQLJ translation program that replaces the embedded SQLJ code with calls to an SQLJ Java library. The modified program code can then be compiled by any Java compiler. Usually the SQLJ Java library makes calls to a JDBC driver, which handles the connection to the database system.

An important philosophical difference exists between Embedded SQL and SQLJ and JDBC. Since vendors provide their own proprietary versions of SQL, it is advisable to write SQL queries according to the SQL-92 or SQL:1999 standard. However, when using Embedded SQL, it is tempting to use vendor-specific SQL constructs that offer functionality beyond the SQL-92 or SQL:1999 standards. SQLJ and JDBC force adherence to the standards, and the resulting code is much more portable across different database systems.

In the remainder of this section, we give a short introduction to SQLJ.

6.4.1 Writing SQLJ Code

We will introduce SQLJ by means of examples. Let us start with an SQLJ code fragment that selects records from the Books table that match a given author.

```
String title; Float price; String atithor;
#sql iterator Books (String title, Float price);
Books books;

// the application sets the author
// execute the query and open the cursor
#sql books = {
    SELECT title, price INTO :titIe, :price
    FROM Books WHERE author = :author
};

// retrieve results
while (books.next()) {
    System.out.println(books.titleO + ", " + books.price());
}
books.close();
```

The corresponding JDBC code fragment looks as follows (assuming we also declared price, name, and author:

```
PrcparcdStatcment stmt = connection.prepareStatement(
"SELECT title, price FROM Books WHERE author = ?");

// set the parameter in the query ancl execute it
stmt.setString(1, author);
ResultSet 1'8 = stmt.executeQuery();

// retrieve the results
while (rs.next()) {
```

```
System.out.println(rs.getString(l) + ", " + rs.getFloat(2));
}
```

Comparing the JDBC and SQLJ code, we see that the SQLJ code is much easier to read than the JDBC code. Thus, SQLJ reduces software development and maintenance costs.

Let us consider the individual components of the SQLJ code in more detail. All SQLJ statements have the special prefix #sql. In SQLJ, we retrieve the results of SQL queries with iterator objects, which are basically cursors. An iterator is an instance of an iterator class. Usage of an iterator in SQLJ goes through five steps:

- Declare the Iterator Class: In the preceding code, this happened through the statement
  #sql iterator Books (String title, Float price);
  This statement creates a new Java class that we can use to instantiate objects.
- Instantiate an Iterator Object from the New Iterator Class: We instantiated our iterator in the statement Books books;.
- Initialize the Iterator Using a SQL Statement: In our example, this happens through the statement #sql books = ...
- Iteratively, Read the Rows From the Iterator Object: This step is very similar to reading rows through a ResultSet object in JDBC.
- Close the Iterator Object.

There are two types of iterator classes: named iterators and positional iterators. For named iterators, we specify both the variable type and the name of each column of the iterator. This allows us to retrieve individual columns by name as in our previous example where we could retrieve the title colunm from the Books table using the expression books.titIe(). For positional iterators, we need to specify only the variable type for each column of the iterator. To access the individual columns of the iterator, we use a FETCH ... INTO eonstruct, similar to Embedded SQL. Both iterator types have the same performance; which iterator to use depends on the programmer's taste.

Let us revisit our example. \Ve can make the iterator a positional iterator through the following statement:

```
#sql iterator Books (String, Float);
```

We then retrieve the individual rows from the iterator as follows:

```
while (true) {
    #sql { FETCH :books INTO :title, :price, };
    if (books.endFetch()) {
        break:
    }

    // process the book
}
```

6.5 STORED PROCEDURES

It is often important to execute some parts of the application logic directly in the process space of the database system. Running application logic directly at the database has the advantage that the amount of data that is transferred between the database server and the client issuing the SQL statement can be minimized, while at the same time utilizing the full power of the database server.

When SQL statements are issued from a remote application, the records in the result of the query need to be transferred from the database system back to the application. If we use a cursor to remotely access the results of an SQL statement, the DBMS has resources such as locks and memory tied up while the application is processing the records retrieved through the cursor. In contrast, a stored procedure is a program that is executed through a single SQL statement that can be locally executed and completed within the process space of the database server. The results can be packaged into one big result and returned to the application, or the application logic can be performed directly at the server, without having to transmit the results to the client at all.

Stored procedures are also beneficial for software engineering reasons. Once a stored procedure is registered with the database server, different users can re-use the stored procedure, eliminating duplication of efforts in writing SQL queries or application logic, and making code maintenance easy. In addition, application programmers do not need to know the the database schema if we encapsulate all database access into stored procedures.

Although they are called stored *procedures*, they do not have to be procedures in a programming language sense; they can be functions.

6.5.1 Creating a Simple Stored Procedure

Let us look at the example stored procedure written in SQL shown in Figure 6.8. We see that stored procedures must have a name; this stored procedure

has the name 'ShowNumberOfOrders.' Otherwise, it just contains an SQL statement that is precompiled and stored at the server.

```
CREATE PROCEDURE ShowNumberOfOrders
SELECT C.cid, C.cname, COUNT(*)
       FROM    Customers C, Orders a
       WHERE   C.cid = O.cid
       GROUP BY C.cid, C.cname
```

Figure 6.8 A Stored Procedure in SQL

Stored procedures can also have parameters. These parameters have to be valid SQL types, and have one of three different modes: IN, OUT, or INOUT. IN parameters are arguments to' the stored procedure. OUT parameters are returned from the stored procedure; it assigns values to all OUT parameters that the user can process. INOUT parameters combine the properties of IN and OUT parameters: They contain values to be passed to the stored procedures, and the stored procedure can set their values as return values. Stored procedures enforce strict type conformance: If a parameter is of type INTEGER, it cannot be called with an argument of type VARCHAR.

Let us look at an example of a stored procedure with arguments. The stored procedure shown in Figure 6.9 has two arguments: book_isbn and addedQty. It updates the available number of copies of a book with the quantity from a new shipment.

```
CREATE PROCEDURE AddInventory (
                 IN book_isbn CHAR(lO),
                 IN addedQty INTEGER)
UPDATE Books
       SET     qty_in_stock = qtyjn_stock + addedQty
       WHERE   bookjsbn = isbn
```

Figure 6.9 A Stored Procedure with Arguments

Stored procedures do not have to be written in SQL; they can be written in any host language. As an example, the stored procedure shown in Figure 6.10 is a Java function that is dynamically executed by the database server whenever it is called by the dient:

6.5.2 Calling Stored Procedures

Stored procedures can be called in interactive SQL with the CALL statement:

```
CREATE PROCEDURE RallkCustomers(IN number INTEGER)
LANGUAGE Java
EXTERNAL NAME 'file:///c:/storedProcedures/rank.jar'
```

Figure 6.10 A Stored Procedure in Java

```
CALL storedProcedureName(argumentl, argument2, ... , argumentN);
```

In Embedded SQL, the arguments to a stored procedure are usually variables in the host language. For example, the stored procedure AddInventory would be called as follows:

```
EXEC SQL BEGIN DECLARE SECTION
char isbn[lO];
long qty;
EXEC SQL END DECLARE SECTION

// set isbn and qty to some values
EXEC SQL CALL AddInventory(:isbn,:qty);
```

Calling Stored Procedures from JDBC

We can call stored procedures from JDBC using the CallableStatment class. CallableStatement is a subclass of PreparedStatement and provides the same functionality. A stored procedure could contain multiple SQL staternents or a series of SQL statements-thus, the result could be many different ResultSet objects. We illustrate the case when the stored procedure result is a single ResultSet.

```
CallableStatement cstmt=
            con.prepareCall("{call ShowNumberOfOrders}");
ResultSet rs = cstmt.executeQueryO
while (rs.next())
```

Calling Stored Procedures from SQLJ

The stored procedure 'ShowNumberOfOrders' is called as follows using SQLJ:

```
// create the cursor class
#sql !terator CustomerInfo(int cid, String cname, int count);

// create the cursor
```

```
CustomerInfo customerinfo;

// call the stored procedure
#sql customerinfo = {CALL ShowNumberOfOrders};
while (customerinfo.nextO) {
    System.out.println(customerinfo.cid() + "," +
                customerinfo.count());
}
```

6.5.3 SQLIPSM

All major database systems provide ways for users to write stored procedures in a simple, general purpose language closely aligned with SQL. In this section, we briefly discuss the SQL/PSM standard, which is representative of most vendor-specific languages. In PSM, we define modules, which are collections of stored procedures, temporary relations, and other declarations.

In SQL/PSM, we declare a stored procedure as follows:

```
CREATE PROCEDURE name (parameter1,..., parameterN)
    local variable declarations
    procedure code;
```

We can declare a function similarly as follows:

```
CREATE FUNCTION name (parameterl,..., parameterN)
            RETURNS sqIDataType
    local variable declarations
    function code;
```

Each parameter is a triple consisting of the mode (IN, OUT, or INOUT as discussed in the previous section), the parameter name, and the SQL datatype of the parameter. We can seen very simple SQL/PSM procedures in Section 6.5.1. In this case, the local variable declarations were empty, and the procedure code consisted of an SQL query.

We start out with an example of a SQL/PSM function that illustrates the main SQL/PSM constructs. The function takes as input a customer identified by her *cid* and a year. The function returns the rating of the customer, which is defined as follows: Customers who have bought more than ten books during the year are rated 'two'; customer who have purchased between 5 and 10 books are rated 'one', otherwise the customer is rated 'zero'. The following SQL/PSM code computes the rating for a given customer and year.

```
CREATE PROCEDURE RateCustomer
```

```
                    (IN custId INTEGER, IN year INTEGER)
                    RETURNS INTEGER
DECLARE rating INTEGER;
DECLARE numOrders INTEGER;
SET numOrders =
        (SELECT COUNT(*) FROM Orders 0 WHERE O.tid = custId);
IF (numOrders> 10) THEN rating=2;
ELSEIF (numOrders>5) THEN rating=1;
ELSE rating=O;
END IF;
RETURN rating;
```

Let us use this example to give a short overview of some SQL/PSM constructs:

- We can declare local variables using the DECLARE statement. In our example, we declare two local variables: 'rating', and 'numOrders'.
- PSM/SQL functions return values via the RETURN statement. In our example, we return the value of the local variable 'rating'.
- We can assign values to variables with the SET statement. In our example, we assigned the return value of a query to the variable 'numOrders'.
- SQL/PSM has branches and loops. Branches have the following form:

```
IF (condition) THEN statements;
ELSEIF statements;

ELSEIF statements;
ELSE statements; END IF
```

  Loops are of the form

```
LOOP
        staternents:
END LOOP
```

- Queries can be used as part of expressions in branches; queries that return a single value can be assigned to variables as in our example above.
- 'We can use the same cursor statements as in Embedded SQL (OPEN, FETCH, CLOSE), but we do not need the EXEC SQL constructs, and variables do not have to be prefixed by a colon ':'.

We only gave a very short overview of SQL/PSM; the references at the end of the chapter provide more information.

6.6 CASE STUDY: THE INTERNET BOOK SHOP

DBDudes finished logical database design, as discussed in Section 3.8, and now consider the queries that they have to support. They expect that the application logic will be implemented in Java, and so they consider JDBC and SQLJ as possible candidates for interfacing the database system with application code.

Recall that DBDudes settled on the following schema:

Books(*isbn:* CHAR(10), *title:* CHAR(8), *author:* CHAR(80),
qty_in_stock: INTEGER, *price:* REAL, *year_published:* INTEGER)
Customers(*cid:* INTEGER, *cname:* CHAR(80), *address:* CHAR(200))
Orders(*ordernum:* INTEGER, *isbn:* CHAR(lO), *cid:* INTEGER,
cardnum: CHAR(l6), *qty:* INTEGER, *order_date:* DATE, *ship_date:* DATE)

Now, DBDudes considers the types of queries and updates that will arise. They first create a list of tasks that will be performed in the application. Tasks performed by customers include the following.

- Customers search books by author name, title, or ISBN.
- Customers register with the website. Registered customers might want to change their contact information. DBDudes realize that they have to augment the Customers table with additional information to capture login and password information for each customer; we do not discuss this aspect any further.
- Customers check out a final shopping basket to complete a sale.
- Customers add and delete books from a 'shopping basket' at the website.
- Customers check the status of existing orders and look at old orders.

Administrative tasks performed by employees of B&N are listed next.

- Employees look up customer contact information.
- Employees add new books to the inventory.
- Employees fulfill orders, and need to update the shipping date of individual books.
- Employees analyze the data to find profitable customers and customers likely to respond to special marketing campaigns.

Next, DBDudes consider the types of queries that will arise out of these tasks. To support searching for books by name, author, title, or ISBN, DBDudes decide to write a stored procedure as follows:

```
CREATE PROCEDURE SearchByISBN (IN book.isbn CHAR(10))
    SELECT B.title, B.author, B.qty_in_stock, B.price, B.yeaLpublished
    FROM   Books B
    WHERE  B.isbn = book.isbn
```

Placing an order involves inserting one or more records into the Orders table. Since DBDudes has not yet chosen the Java-based technology to program the application logic, they assume for now that the individual books in the order are stored at the application layer in a Java array. To finalize the order, they write the following JDBC code shown in Figure 6.11, which inserts the elements from the array into the Orders table. Note that this code fragment assumes several Java variables have been set beforehand.

```
String sql = "INSERT INTO Orders VALUES(7, 7, 7, 7, 7, 7)";
PreparedStatement pstmt = con.prepareStatement(sql);
con.setAutoCommit(false);

try {
    // orderList is a vector of Order objects
    // ordernum is the current order number
    // dd is the ID of the customer, cardnum is the credit card number
    for (int i=O; iiorderList.lengthO; i++)
        // now instantiate the parameters with values
        Order currentOrder = orderList[i];
        pstmt.clearParameters();
        pstmt.setInt(l, ordernum);
        pstmt.setString(2, Order.getlsbnO);
        pstmt.setInt(3, dd);
        pstmt.setString(4, creditCardNum);
        pstmt.setlnt(5, Order.getQtyO);
        pstmt.setDate(6, null);

        pstmt.executeUpdate();
    }
    con.commit();
catch (SqLException e){
    con.rollbackO;
    System.out.println(e.getMessage());
}
```

Figure 6.11 Inserting a Completed Order into the Database

DBDudes writes other JDBC code and stored procedures for all of the remaining tasks. They use code similar to some of the fragments that we have seen in this chapter.

- Establishing a connection to a database, as shown in Figure 6.2.
- Adding new books to the inventory, as shown in Figure 6.3.
- Processing results from SQL queries as shown in Figure 6.4-
- For each customer, showing how many orders he or she has placed. We showed a sample stored procedure for this query in Figure 6.8.
- Increasing the available number of copies of a book by adding inventory, as shown in Figure 6.9.
- Ranking customers according to their purchases, as shown in Figure 6.10.

DBDudcs takes care to make the application robust by processing exceptions and warnings, as shown in Figure 6.6.

DBDudes also decide to write a trigger, which is shown in Figure 6.12. Whenever a new order is entered into the Orders table, it is inserted with ship_date set to NULL. The trigger processes each row in the order and calls the stored procedure 'UpdateShipDate'. This stored procedure (whose code is not shown here) updates the (anticipated) ship_date of the new order to 'tomorrow', in case qtyjlLstock of the corresponding book in the Books table is greater than zero. Otherwise, the stored procedme sets the ship_date to two weeks.

```
CREATE TRIGGER update_ShipDate
        AFTER INSERT ON Orders                    /* Event */
  FOR EACH ROW
  BEGIN CALL UpdatcShipDate(new); END             /* Action */
```

Figure 6.12 Trigger to Update the Shipping Date of New Orders

6.7 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Why is it not straightforward to integrate SQL queries with a host programming language? (Section 6.1.1)
- How do we declare variables in Ernbcdded SQL? (Section 6.1.1)

- How do we use SQL statements within a host langl.lage? How do we check for errors in statement execution? (Section 6.1.1)
- Explain the impedance mismatch between host languages and SQL, and describe how cursors address this. (Section 6.1.2)
- What properties can cursors have? (Section 6.1.2)
- What is Dynamic SQL and how is it different from Embedded SQL? (Section 6.1.3)
- What is JDBC and what are its advantages? (Section 6.2)
- What are the components of the JDBC architecture? Describe four different architectural alternatives for JDBC drivers. (Section 6.2.1)
- How do we load JDBC drivers in Java code? (Section 6.3.1)
- How do we manage connections to data sources? What properties can connections have? (Section 6.3.2)
- What alternatives does JDBC provide for executing SQL DML and DDL statements? (Section 6.3.3)
- How do we handle exceptions and warnings in JDBC? (Section 6.3.5)
- 'What functionality provides the DatabaseMetaDataclass? (Section 6.3.6)
- What is SQLJ and how is it different from JDBC? (Section 6.4)
- Why are stored procedures important? How do we declare stored procedures and how are they called from application code? (Section 6.5)

EXERCISES

Exercise 6.1 Briefly answer the following questions.

1. Explain the following terms: Cursor, Embedded SQL, JDBC, SQLJ, stored procedure.
2. What are the differences between JDBC and SQLJ? \Nhy do they both exist?
3. Explain the term *stored procedure*, and give examples why stored procedures are useful.

Exercise 6.2 Explain how the following steps are performed in JDBC:

1. Connect to a data source.
2. Start, commit, and abort transactions.
3. Call a stored procedure.

How are these steps performed in SQLJ?

Exercise 6.3 Compare exception handling and handling of warnings ill embedded SQL, dynamic SQL, .IDBC, and SQL.I.

Exercise 6.4 Answer the following questions.

1. Why do we need a precompiler to translate embedded SQL and SQL.J? Why do we not need a precompiler for .IDBC?
2. SQL.J and embedded SQL use variables in the host language to pass parameters to SQL queries, whereas .JDBC uses placeholders marked with a '?'. Explain the difference, and why the different mechanisms are needed.

Exercise 6.5 A dynamic web site generates HTML pages from information stored in a database. Whenever a page is requested, is it dynamically assembled from static data and data in a database, resulting in a database access. Connecting to the database is usually a time-consuming process, since resources need to be allocated, and the user needs to be authenticated. Therefore, connection pooling--setting up a pool of persistent database connections and then reusing them for different requests can significantly improve the performance of database-backed websites. Since servlets can keep information beyond single requests, we can create a connection pool, and allocate resources from it to new requests.

Write a connection pool class that provides the following methods:

- Create the pool with a specified number of open connections to the database system.
- Obtain an open connection from the pool.
- Release a connection to the pool.
- Destroy the pool and close all connections.

PROJECT-BASED EXERCISES

In the following exercises, you will create database-backed applications. In this chapter, you will create the parts of the application that access the database. In the next chapter, you will extend this code to other aspects of the application. Detailed information about these exercises and material for more exercises can be found online at

http://www.cs.wisc.edu/-dbbook

Exercise 6.6 Recall the Notown Records database that you worked with in Exercise 2.5 and Exercise 3.15. You have now been tasked with designing a website for Notown. It should provide the following functionality:

- Users can search for records by name of the musician, title of the album, and Bame of the song.
- Users can register with the site, and registered users ca.n log on to the site. Once logged on, users should not have to log on again unless they are inactive for a long time.
- Users who have logged on to the site can add items to a shopping basket.
- Users with items in their shopping basket can check out and ma.ke a purchase.

NOtOWll wants to use JDBC to access the database. Write .JDBC code that performs the necessary data access and manipulation. You will integrate this code with application logic and presentation in the next chapter.

If Notown had chosen SQLJ instead of JDBC, how would your code change?

Exercise 6.7 Recall the database schema for Prescriptions-R-X that you created in Exercise 2.7. The Prescriptions-R-X chain of pharmacies has now engaged you to design their new website. The website has two different classes of users: doctors and patients. Doctors should be able to enter new prescriptions for their patients and modify existing prescriptions. Patients should be able to declare themselves as patients of a doctor; they should be able to check the status of their prescriptions online; and they should be able to purchase the prescriptions online so that the drugs can be shipped to their home address.

Follow the analogous steps from Exercise 6.6 to write JDBC code that performs the necessary data access and manipulation. You will integrate this code with application logic and presentation in the next chapter.

Exercise 6.8 Recall the university database schema that you worked with in Exercise 5.l. The university has decided to move enrollment to an online system. The website has two different classes of users: faculty and students. Faculty should be able to create new courses and delete existing courses, and students should be able to enroll in existing courses.

Follow the analogous steps from Exercise 6.6 to write JDBC code that performs the necessary data access and manipulation. You will integrate this code with application logic and presentation in the next chapter.

Exercise 6.9 Recall the airline reservation schema that you worked on in Exercise 5.3. Design an online airline reservation system. The reservation system will have two types of users: airline employees, and airline passengers. Airline employees can schedule new flights and cancel existing flights. Airline passengers can book existing flights from a given destination.

Follow the analogous steps from Exercise 6.6 to write JDBC code that performs the necessary data access and manipulation. You will integrate this code with application logic and presentation in the next chapter.

BIBLIOGRAPHIC NOTES

Information on ODBC can be found on Microsoft's web page (www.microsoft.com/data/odbc), and information on JDBC can be found on the Java web page (java. sun. com/products/jdbc). There exist rnany books on ODBC, for example, Sanders' ODBC Developer's Guicle [652] and the lvIicrosoft ODBC SDK [533]. Books on JDBC include works by Hamilton et al. [359], Reese [621], and White et a!. [773].

7

INTERNET APPLICATIONS

- How do we name resources on the Internet?
- How do Web browsers and webservers communicate?
- How do we present documents on the Internet? How do we differentiate between formatting and content?
- What is a three-tier application architecture? How do we write three-tiered applications?
- Why do we have application servers?
- Key concepts: Uniform Resource Identifiers (URI), Uniform Resource Locators (URL); Hypertext Transfer Protocol (HTTP), stateless protocol; Java; HTML; XML, XML DTD; three-tier architecture, client-server architecture; HTML forms; JavaScript; cascading style sheets, XSL; application server; Common Gateway Interface (CGI); servlet; JavaServer Page (JSP); cookie

Wow! They've got the Internet on computers now!

--Homer Simpson, *The Simpsons*

7.1 INTROpUCTION

The proliferation of computer networks, including the Internet and corporate 'intranets,' has enabled users to access a large number of data sources. This increased access to databases is likely to have a great practical impact; data and services can now be offered directly to customers in ways impossible until

recently. Examples of such electronic commerce applications include purchasing books through a \Veb retailer such as Amazon.com, engaging in online auctions at a site such as eBay, and exchanging bids and specifications for products between companies. The emergence of standards such as XrvIL for describing the content of documents is likely to further accelerate electronic commerce and other online applications.

While the first generation of Internet sites were collections of HTML files, most major sites today store a large part (if not all) of their data in database systems. They rely on DBMSs to provide fast, reliable responses to user requests received over the Internet. This is especially true of sites for electronic commerce and other business applications.

In this chapter, we present an overview of concepts that are central to Internet application development. We start out with a basic overview of how the Internet works in Section 7.2. We introduce HTML and XML, two data formats that are used to present data on the Internet, in Sections 7.3 and 7.4. In Section 7.5, we introduce three-tier architectures, a way of structuring Internet applications into different layers that encapsulate different functionality. In Sections 7.6 and 7.7, we describe the presentation layer and the middle layer in detail; the DBMS is the third layer. We conclude the chapter by discussing our B&N case study in Section 7.8.

Examples that appear in this chapter are available online at

http://www.cs.wisc.edu/-dbbook

7.2 INTERNET CONCEPTS

The Internet has emerged as a universal connector between globally distributed software systems. To understand how it works, we begin by discussing two basic issues: how sites on the Internet are identified, and how programs at one site communicate with other sites.

We first introduce Uniform Resource Identifiers, a naming schema for locating resources on the Internet in Section 7.2.1. \Ve then talk about the most popular protocol for accessing resources over the Web, the hypertext transfer protocol (HTTP) in Section 7.2.2.

7.2.1 Uniform Resource Identifiers

Uniform Resource Identifiers (URIs), are strings that uniquely identify resources 011 the Internet. A resource is any kind of information that can

Distributed Applications and Service-Oriented Architectures:

The advent of XML, due to its loosely-coupled nature, has made information exchange between different applications feasible to an extent previously unseen. By using XML for information exchange, applications can be written in different programming languages, run on different operating systems, and yet they can still share information with each other. There are also standards for externally describing the intended content of an XML file or message, most notably the recently adopted W3C XML Schemas standard.

A promising concept that has arisen out of the XML revolution is the notion of a Web service. A Web service is an application that provides a well-defined service, packaged as a set of remotely callable procedures accessible through the Internet. Web services have the potential to enable powerful new applications by *composing* existing Web services-all communicating seamlessly thanks to the use of standardizedXML-based information exchange. Several technologies have been developed or are currently under development that facilitate design and implementation of distributed applications. SOAP is a W3C standard for XML-based invocation of remote services (think XML RPC) that allows distributed applications to communicate either synchronously or asynchronously via structured, typed XML messages. SOAP calls can ride on a variety of underlying transport layers, including HTTP (part of what is making SOAP so successful) and various reliable messaging layers. Related to the SOAP standard are W3C's Web Services Description Language (WSDL) for describing Web service interfaces, and Universal Description, Discovery, and Integration (UDDI), a WSDL-based Web services registry standard (think yellow pages for Web services).

SOAP-based Web services are the foundation for Microsoft's recently released .NET framework, their application development infrastructure and associated run-time system for developing distributed applications, as well as for the Web services offerings of major software vendors such as IBM, BEA, and others. Many large software application vendors (major companies like PeopleSoft and SAP) have announced plans to provide Web service interfaces to their products and the data that they manage, and many are hoping that XML and Web services will finally provide the answer to the long-standing problem of enterprise application integration. Web services are also being looked to as a natural foundation for the next generation of business process management (or workflow) systems.

be identified by a URI, and examples include webpages, images, downloadable files, services that can be remotely invoked, mailboxes, and so on. The most common kind of resource is a static file (such as a HTML document), but a resource may also be a dynamically-generated HTML file, a movie, the output of a program, etc.

A URI has three parts:

- The (name of the) protocol used to access the resource.
- The host computer where the resource is located.
- The path name of the resource itself on the host computer.

Consider an example URI, such as http://www.bookstore.com/index.html. This URI can be interpreted as follows. Use the HTTP protocol (explained in the next section) to retrieve the document index.html located at the computer www.bookstore.com.This example URI is an instance of a Universal Resource Locator (URL), a subset of the more general URI naming scheme; the distinction is not important for our purposes. As another example, the following HTML fragment shows a URI that is an email address:

```
<a href=lImailto:webmaster@bookstore.com">Email the webmaster.</A>
```

7.2.2 The Hypertext Transfer Protocol (HTTP)

A communication protocol is a set of standards that defines the structure of messages between two communicating parties so that they can understand each other's messages. The Hypertext Transfer Protocol (HTTP) is the most common communication protocol used over the Internet. It is a client-server protocol in which a client (usually a Web browser) sends a request to an HTTP server, which sends a response back to the client. When a user requests a webpage (e.g., clicks on a hyperlink), the browser sends **HTTP** request messages for the objects in the page to the server. The server receives the requests and responds with **HTTP** response messages, which include the objects. It is important to recognize that HTTP is used to transmit all kinds of resources, not just files, but most resources on the Internet today are either static files or files output from server-side scripts.

A variant of the HTTP protocol called the Secure Sockets Layer (SSL) protocol uses encryption to exchange information securely between client and server. We postpone a discussion of SSL to Section 21.5.2 and present the basic HTTP protocol in this chapter.

As an example, consider what happens if a user clicks on the following link: http://www.bookstore.com/index.html. We first explain the structure of an HTTP request message and then the structure of an HTTP response message.

HTTP Requests

The client (Web browser) establishes a connection with the webserver that hosts the resource and sends a HTTP request message. The following example shows a sample HTTP request message:

```
GET index.html HTTP/1.1
User-agent: Mozilla/4.0
Accept: text/html, image/gif, image/jpeg
```

The general structure of an HTTP request consists of several lines of ASCII text, with an empty line at the end. The first line, the request line, has three fields: the HTTP method field, the URI field, and the HTTP version field. The method field can take on values GET and POST; in the example the message requests the object index.html. (We discuss the differences between HTTP GET and HTTP POST in detail in Section 7.11.) The version field indicates which version of HTTP is used by the client and can be used for future extensions of the protocol. The user agent indicates the type of the client (e.g., versions of Netscape or Internet Explorer); we do not discuss this option further. The third line, starting with Accept, indicates what types of files the client is willing to accept. For example, if the page index.html contains a movie file with the extension .mpg, the server will not send this file to the client, as the client is not ready to accept it.

HTTP Responses

The server responds with an HTTP response message. It retrieves the page index.html, uses it to assemble the HTTP response message, and sends the message to the client. A sample HTTP response looks like this:

```
HTTP/1.1 200 OK
Date: Mon, 04 Mar 2002 12:00:00 GMT
Content-Length: 1024
Content-Type: text/html
Last-Modified: Mall, 22 Jun 1998 09:23:24 GMT
<HTML>
<HEAD>
</HEAD>
<BODY>
```

<H1>Barns and Nobble Internet Bookstore</H1>
Our inventory:
<H3>Science</H3>
<B>The Character of Physical Law</B>

The HTTP response message has three parts: a status line, several header lines, and the body of the message (which contains the actual object that the client requested). The status line has three fields (analogous to the request line of the HTTP request message): the HTTP version (HTTP/1.1), a status code (200), and an associated server message (OK). Common status codes and associated messages are:

- 200 OK: The request succeeded and the object is contained in the body of the response message";
- 400 Bad Request: A generic error code indicating that the request could not be fulfilled by the server.
- 404 Not Found: The requested object does not exist on the server.
- 505 HTTP Version Not Supported: The HTTP protocol version that the client uses is not supported by the server. (Recall that the HTTP protocol version sent in the client's request.)

Our example has three header lines: The date header line indicates the time and date when the HTTP response was created (not that this is not the object creation time). The Last-Modified header line indicates when the object was created. The Content-Length header line indicates the number of bytes in the object being sent after the last header line. The Content-Type header line indicates that the object in the entity body is HTML text.

The client (the Web browser) receives the response message, extracts the HTML file, parses it, and displays it. In doing so, it might find additional URIs in the file, and it then uses the HTTP protocol to retrieve each of these resources, establishing a new connection each time.

One important issue is that the HTTP protocol is a stateless protocol. Every message----from, the client to the HTTP server and vice-versa-is self-contained, and the connection established with a request is maintained only until the response message is sent. The protocol provides no mechanism to automatically 'remember' previous interactions between client and server.

The stateless nature of the HTTP protocol has a major impact on how Internet applications are written. Consider a user who interacts with our exalIlple

bookstore application. Assume that the bookstore permits users to log into the site and then carry out several actions, such as ordering books or changing their address, without logging in again (until the login expires or the user logs out). How do we keep track of whether a user is logged in or not? Since HTTP is stateless, we cannot switch to a different state (say the 'logged in' state) at the protocol level. Instead, for every request that the user (more precisely, his or her Web browser) sends to the server, we must encode any *state* information required by the application, such as the user's login status. Alternatively, the server-side application code must maintain this state information and look it up on a per-request basis. This issue is explored further in Section 7.7.5.

Note that the statelessness of HTTP is a tradeoff between ease of implementation of the HTTP protocol and ease of application development. The designers of HTTP chose to keep the protocol itself simple, and deferred any functionality beyond the request of objects to application layers above the HTTP protocol.

7.3 HTML DOCUMENTS

In this section and the next, we focus on introducing HTML and XML. In Section 7.6, we consider how applications can use HTML and XML to create forms that capture user input, communicate with an HTTP server, and convert the results produced by the data management layer into one of these formats.

HTML is a simple language used to describe a document. It is also called a **markup language** because HTML works by augmenting regular text with 'marks' that hold special meaning for a Web browser. Commands in the language, called tags, consist (usually) of a start tag and an **end** tag of the form <TAG> and </TAG>, respectively. For example, consider the HTML fragment shown in Figure 7.1. It describes a webpage that shows a list of books. The document is enclosed by the tags <HTML> and </HTML>, marking it as an HTML document. The remainder of the document-enclosed in <BODY> ... </BoDY>-contains information about three books. Data about each book is represented as an unordered list (UL) whose entries are marked with the LI tag. HTML defines the set of valid tags as well as the meaning of the tags. For example, HTML specifies that the tag <TITLE> is a valid tag that denotes the title of the document. As another example, the tag <UL> always denotes an unordered list.

Audio, video, and even programs (written in Java, a highly portable language) can be included in HTML documents. When a user retrieves such a document using a suitable browser, images in the document arc displayed, audio and video clips are played, and embedded programs are executed at the user's machine; the result is a rich multimedia presentation. The case with which HTML docu-

```
<HTML>
<HEAD>
</HEAD>
<BODY>
<Hl>Barns and Nobble Internet Bookstore</Hl>
Our inventory:
<H3>Science</H3>
    <B>The Character of Physical Law</B>
    <UL>
        <LI>Author: Richard Feynman</LI>
        <LI>Published 1980</LI>
        <Ll>Hardcover</LI>
    </UL>
<H3>Fiction</H3>
    <B>Waiting for the Mahatma</B>
    <UL>
        <LI>Author: R.K. Narayan</LI>
        <LI>Published 1981</Ll>
    </UL>
    <B>The English Teacher</B>
    <UL>
        <LI>Author: R.K. Narayan</LI>
        <LI>Published 1980</LI>
        <LI>Paperback</LI>
    </UL>
</BODY>
</HTML>
```

Figure 7.1 Book Listing in HTML

ments can be created—there are now visual editors that automatically generate HTML----and accessed using Internet browsers has fueled the explosive growth of the Web.

7.4 XML DOCUMENTS

In this section, we introduce XML as a document format, and consider how applications can utilize XML. Managing XML documents in a DBMS poses several new challenges; we discuss this aspect of XML in Chapter 27.

vVhile HTML can be used to mark up documents for display purposes, it is not adequate to describe the structure of the content for more general applications. For example, we can send the HTML document shown in Figure 7.1 to another application that displays it, but the second application cannot distinguish the first names of authors from their last names. (The application can try to recover such information by looking at the text inside the tags, but this defeats the purpose of using tags to describe document structure.) Therefore, HTML is unsuitable for the exchange of complex documents containing product specifications or bids, for example.

Extensible Markup Language (XML) is a markup language developed to remedy the shortcomings of HTML. In contrast to a fixed set of tags whose meaning is specified by the language (as in HTML), XML allows users to define new collections of tags that can be used to structure any type of data or document the user wishes to transmit. XML is an important bridge between the document-oriented view of data implicit in HTML and the schema-oriented view of data that is central to a DBMS. It has the potential to make database systems more tightly integrated into Web applications than ever before.

XML emerged from the confluence of two technologies, SGML and HTML. The Standard Generalized Markup Language (SGML) is a metalanguage that allows the definition of data and document interchange languages such as HTML. The SGML standard was published in 1988, and many organizations that rnanage a large number of complex documents have adopted it. Due to its generality, SGML is complex and requires sophisticated programs to harness its full potential. XML was developed to have much of the power of SGML while remaining relatively simple. Nonetheless, XML, like SGML, allows the definition of new document markup languages.

Although XML does not prevent a user from designing tags that encode the display of the data in a Web browser, there is a style language for XML called Extensible Style Language (XSL). XSL is a standard way of describing how an XML docmnent that adheres to a certain vocabulary of tags should be displayed.

7.4.1 Introduction to XML

We use the small XML docmnent shown in Figure 7.2 as an example.

- Elements: Elements, also called tags, arc the primary building blocks of an XML docmnent. The start of the content of an element ELM is marked with <ELM>, which is called the start tag, and the end of the content end is marked with </ELM>, called the end tag. In our example document.

The Design Goals of XML: XML was developed starting in 1996 by a working group under guidance of the World Wide Web Consortium (W3C) XML Special Interest Group. The design goals for XML included the following:

1. XML should be compatible with SGML.
2. It should be easy to write programs that process XML documents.
3. The design of XML should be formal and concise.

the element BOOKLIST encloses all information in the sample document. The element BOOK demarcates all data associated with a single book. XML elements are case sensitive: the element BOOK is different from Book. Elements must be properly nested. Start tags that appear inside the content of other tags must have a corresponding end tag. For example, consider the following XML fragment:

```
<BOOK>
    <AUTHOR>
        <FIRSTNAME>Richard</FIRSTNAME>
        <LASTNAME>Feynluan</LASTNAME>
    </AUTHOR>
</BOOK>
```

The element AUTHOR is completely nested inside the element BOOK, and both the elements LASTNAME and FIRSTNAME are nested inside the element AUTHOR.

- Attributes: An element can have descriptive attributes that provide additional information about the element. The values of attributes are set inside the start tag of an element. For example, let ELM denote an element with the attribute att. We can set the value of att to value through the following expression: <ELM att="value">. All attribute values must be enclosed in quotes. In Figure 7.2, the element BOOK has two attributes. The attribute GENRE indicates the genre of the book (science or fiction) and the attribute FORMAT indicates whether the book is a hardcover or a paperback.

- Entity References: Entities are shortcuts for portions of common text or the content of external files, and we call the usage of an entity in the XML document an entity reference. Wherever an entity reference appears in the document, it is textually replaced by its content. Entity references start with a '&' and end with a ';'. Five predefined entities in XML are placeholders for chara.cters with special meaning in XML. For example, the

```
<?xml version=11.0" encoding="UTF-S" standalone=llyesll?>
<BOOKLIST>
<BOOK GENRE=" Science" FORMAT=" Hardcover" >
    <AUTHOR>
        <FIRSTNAME>Richard</FIRSTNAME>
        <LASTNAME>Feynman</LASTNAME>
    </AUTHOR>
    <TITLE>The Character of Physical Law</TITLE>
    <PUBLISHED>1980</PUBLISHED>
</BOOK>
<BOOK> GENRE=" Fiction" >
    <AUTHOR>
        <FIRSTNAME>R.K.</FIRSTNAME>
        <LASTNAME>Narayan</LASTNAME>
    </AUTHOR>
    <TITLE>Waiting for the Mahatma</TITLE>
    <PUBLISHED>1981</PUBLISHED>
</BOOK>
<BOOK GENRE=" Fiction" >
    <AUTHOR>
        <FIRSTNAME>R.K.</FIRSTNAME>
        <LASTNAME>Narayan</LASTNAME>
    </AUTHOR>
    <TITLE>The English Teacher</TITLE>
    <PUBLISHED> 1980</PUBLISHED>
</BOOK>
</BOOKLIST>
```

Figure 7.2 Book Information in XML

< character that marks the beginning of an XML command is reserved and has to be represented by the entity It. The other four reserved characters are &, >, ", and '; they are represented by the entities amp, gt, quot, and apos. For example, the text '1 < 5' has to be encoded in an XML document as follows: ' 1&1t ;5'. We can also use entities to insert arbitrary Unicode characters into the text. Unicode is a standard for character representations, similar to ASCII. For example, we can display the Japanese Hiragana character *a* using the entity reference &#x3042.

- Comments: We can insert comments anywhere in an XML document. Comments start with <! - and end with ->. Comments can contain arbitrary text except the string --.

- Document Type Declarations (DTDs): In XML, we can define our own markup language. A DTD is a set of rules that allows us to specify our own set of elements, attributes, and entities. Thus, a DTD is basically a grammar that indicates what tags are allowed, in what order they can appear, and how they can be nested. We discuss DTDs in detail in the next section.

We call an XML document well-formed if it has no associated DTD but follows these structural guidelines:

- The document starts with an XML declaration. An example of an XML declaration is the first line of the XML document shown in Figure 7.2.
- A root element contains all the other elements. In our example, the root element is the element BOOKLIST.
- All elements must be properly nested. This requirement states that start and end tags of an element must appear within the same enclosing element.

7.4.2 XML DTDs

A DTD is a set of rules that allows us to specify our own set of elements, attributes, and entities. A DTD specifies which elements we can use and constraints on these elements, for example, how elements can be nested and where elements can appear in the document. We call a document valid if a DTD is associated with it and the document is structured according to the rules set by the DTD. In the remainder of this section, we use the example DTD shown in Figure 7.3 to illustrate how to construct DTDs.

```
<!DOCTYPE BOOKLIST [
<! ELEMENT BOOKLIST (BOOK)*>
    <! ELEMENT BOOK (AUTHOR,TITLE,PUBLISHED?»
        <!ELEMENT AUTHOR (FIRSTNAME,LASTNAME»
            <! ELEMENT FIRSTNAME (#PCDATA»
            <! ELEMENT LASTNAME (#PCDATA»
        <! ELEMENT TITLE (#PCDATA»
        <!ELEMENT PUBLISHED (#PCDATA»
    <! ATTLIST BOOK GENRE (ScienceIFiction) #REQUIRED>
    <!ATTLIST BOOK FORMAT (PaperbackIHardcover) "Paperback">
]>
```

Figure 7.3 Bookstore XML DTD

A DTD is enclosed in <!DOCTYPE name [DTDdeclarationJ >, where name is the name of the outermost enclosing tag, and DTDdeclaration is the text of the rules of the DTD. The DTD starts with the outermost element—the *root element*—which is BOOKLIST in our example. Consider the next rule:

```
<!ELEMENT BOOKLIST (BOOK)*>
```

This rule tells us that the element BOOKLIST consists of zero or more BOOK elements. The * after BOOK indicates how many BOOK elements can appear inside the BOOKLIST element. A * denotes zero or more occurrences, a + denotes one or more occurrences, and a? denotes zero or one occurrence. For example, if we want to ensure that a BOOKLIST has at least one book, we could change the rule as follows:

```
<!ELEMENT BOOKLIST (BOOK)+>
```

Let us look at the next rule:

```
<!ELEMENT BOOK (AUTHOR,TITLE,PUBLISHED?»
```

This rule states that a BOOK element contains a AUTHOR element, a TITLE element, and an optional PUBLISHED clement. Note the use of the? to indicate that the information is optional by having zero or one occurrence of the element. Let us move ahead to the following rule:

```
< !ELEMENT LASTNAME (#PCDATA»
```

Until now we considered only elements that contained other elements. This rule states that LASTNAME is an element that does not contain other elements, but contains actual text. Elements that only contain other elements are said to have element content, whereas elements that also contain #PCDATA are said to have mixed content. In general, an element type declaration has the following structure:

```
<!ELEMENT (contentType»
```

Five possible content types are:

- Other elements.
- The special syrnbol #PCDATA, which indicates (parsed) character data.
- The special symbol EMPTY, which indicates that the element has no content. Elements that have no content are not required to have an end tag.
- The special symbol ANY, which indicates that any content is permitted. This content should be avoided whenever possible since it disables all checking of the document structure inside the element.

- A regular expression constructed from the preceding four choices. A regular expression is one of the following:
  - exp1, exp2, exp3: A list of regular expressions.
  - exp*: An optional expression (zero or more occurrences).
  - exp?: An optional expression (zero or one occurrences).
  - exp+: A mandatory expression (one or more occurrences).
  - expl | exp2: expl or exp2.

Attributes of elements are declared outside the element. For example, consider the following attribute declaration from Figure 7.3:

```
<! ATTLIST BOOK GENRE (ScienceIFiction) #REQUIRED»
```

This XML DTD fragment specifies the attribute GENRE, which is an attribute of the element BOOK. The attribute can take two values: Science or Fiction. Each BOOK element must be described in its start tag by a GENRE attribute since the attribute is required as indicated by #REQUIRED. Let us look at the general structure of a DTD attribute declaration:

```
<! ATTLIST elementName (attName attType default)+>
```

The keyword ATTLIST indicates the beginning of an attribute declaration. The string elementName is the name of the element with which the following attribute dcfinition is associated. What follows is the declaration of one or more attributes. Each attribute has a name, as indicated by attName, and a type, as indicated by attType. XML defines several possible types for an attribute. We discuss only string types and enumerated types here. An attribute of type string can take any string as a value. We can declare such an attribute by setting its type field to CDATA. F'or example, we can declare a third attribute of type string of the elernent BOOK as follows:

```
<!ATTLIST BOOK edition CDATA "1">
```

If an attribute has an enumerated type, we list all its possible values in the attribute declaration. In our example, the attribute GENRE is an enumerated attribute type; its possible attribute values are 'Science' and 'Fiction'.

The last part of an attribute declaration is called its default specification. The DTD in Figure 7.3 shows two different default specifications: #REQUIRED and the string 'Paperback'. The default specification #REQUIRED indicates that the attribute is required and whenever its associated element appears somewhere in the XML document a value for the attribute must be specified. The default specification indicated by the string 'Paperback' indicates that the attribute is not required; whenever its associated element appears without setting

```
<?xml version=11.0" encoding=IUTF-8" standalone="no"?>
<!DOCTYPE BOOKLIST SYSTEM" books.dtd" >
<BOOKLIST>
    <BOOK GENRE=" Science" FORMAT=" Hardcover" >
        <AUTHOR>
```

Figure 7.4 Book Information in XML

XML Schema: The DTD mechanism has several limitations, in spite of its widespread use. For example, elements and attributes cannot be assigned types in a flexible way, and elements are always ordered, even if the application does not require this. XML Schema is a new W3C proposal that provides a more powerful way to describe document structure than DTDs; it is a superset of DTDs, allowing legacy data to be handled easily. An interesting aspect is that it supports uniqueness and foreign key constraints.

a value for the attribute, the attribute automatically takes the value 'Paperback'. For example, we can make the attribute value 'Science' the default value for the GENRE attribute as follows:

```
<! ATTLIST BOOK GENRE (ScienceIFiction) "Science" >
```

In our bookstore example, the XML document with a reference to the DTD is shown in Figure 7.4.

7.4.3 Domain-Specific DTDs

Recently, DTDs have been developed for several specialized domains-including a wide range of commercial, engineering, financial, industrial, and scientific domains----and a lot of the excitement about XML has its origins in the belief that more and more standardized DTDs will be developed. Standardized DTDs would enable seamless data exchange among heterogeneous sources, a problem solved today either by implementing specialized protocols such as Electronic Data Interchange (EDI) or by implementing ad hoc solutions.

Even in an environment where all XML data is valid, it is not possible to straightforwardly integrate several XML documents by matching elements in their DTDs, because even when two elements have identical names in two different DTDs, the meaning of the elements could be completely different. If both documents use a single, standard DTD, we avoid this problem. The

development of standardized DTDs is more a social process than a research problem, since the major players in a given domain or industry segment have to collaborate.

For example, the mathematical markup language (MathML) has been developed for encoding mathematical material on the Web. There are two types of MathML elements. The 28 presentation elements describe the layout structure of a document; examples are the **mrow** element, which indicates a horizontal row of characters, and the msup element, which indicates a base and a subscript. The 75 content elements describe mathematical concepts. An example is the plus element, which denotes the addition operator. (A third type of element, the math element, is used to pass parameters to the MathML processor.) MathML allows us to encode mathematical objects in both notations since the requirements of the user of the objects might be different. Content elements encode the precise mathematical meaning of an object without ambiguity, and the description can be used by applications such as computer algebra systems. On the other hand, good notation can suggest the logical structure to a human and emphasize key aspects of an object; presentation elements allow us to describe mathematical objects at this level.

For example, consider the following simple equation:

$$x^2 - 4x - 32 = 0$$

Using presentation elements, the equation is represented as follows:

```
<mrow>
    <mrow> <msup><mi>x</mi><mn>2</mn></msup>
        <mo>-</mo>
        <mrow><mn>4</mn>
            <mo>&invisibletimes;</mo>
            <mi>x</mi>
        </mrow>
        <mo>-</mo><mn>32</mn>
    </mrow><mo>=</mo><mn>O</nm>
</mrow>
```

Using content elements, the equation is described as follows:

```
<reln><eq/>
    <apply>
        <minus/>
        <apply> <power/> <ci>x</ci> <cn>2</cn> </apply>
        <apply> <times/> <cn>4</cn> <ci>x</ci> </apply>
        <cn>32</cn>
```

```
        </apply> <cn>O</cn>
    </reln>
```

Note the additional power that we gain from using MathML instead of encoding the formula in HTML. The common way of displaying mathematical objects inside an HTML object is to include images that display the objects, for example, as in the following code fragment:

```
<IMG SRC=lIimages/equation.gifu ALI=II x**2 - 4x - 32 = 10 ll >
```

The equation is encoded inside an IMG tag with an alternative display format specified in the ALI tag. Using this encoding of a mathematical object leads to the following presentation problems. First, the image is usually sized to match a certain font size, and on systems with other font sizes the image is either too small or too large. Second, on systems with a different background color, the picture does not blend into the background and the resolution of the image is usually inferior when printing the document. Apart from problems with changing presentations, we cannot easily search for a formula or formula fragments on a page, since there is no specific markup tag.

7.5 THE THREE-TIER APPLICATION ARCHITECTURE

In this section, we discuss the overall architecture of data-intensive Internet applications. Data-intensive Internet applications can be understood in terms of three different functional components: *data management, application logic,* and *pTesentation.* The component that handles data mallgement usually utilizes a DBMS for data storage, but application logic and presentation involve much more than just the DBMS itself.

We start with a short overview of the history of database-backed application architectures, and introduce single-tier and client-server architectures in Section 7.5.1. We explain the three-tier architecture in detail in Section 7.5.2, and show its advantages in Section 7.5.3.

7.5.1 Single-Tier and Client-Server Architectures

In this section, we provide some perspective on the three-tier architecture by discussing single-tier and client-server architectures, the predecessors of the three-tier architecture. Initially, data-intensive applications were combined into a single tier, including the DBMS, application logic, and user interface, as illustrated in Figure 7.5. The application typically ran on a mainframe, and users accessed it through *dumb teT'minals* that could perform only data input and display. This approach has the benefit of being easily maintained by a central administrator.

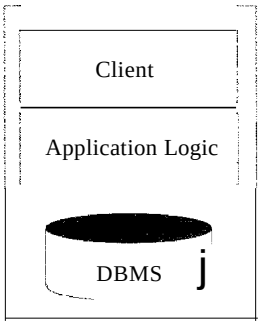


Figure 7.5 A Single-Tier Architecture

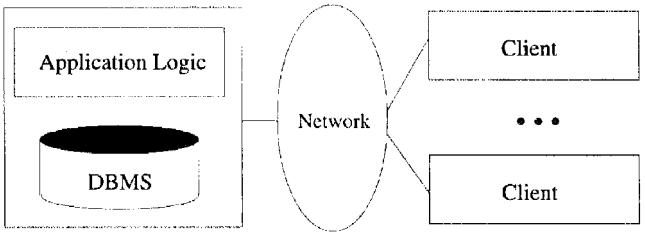


Figure 7.6 A Two-Server Architecture: Thin Clients

Single-tier architectures have a,n important drawback: Users expect graphical interfaces that require much more computational power than simple dumb terminals. Centralized computation of the graphical displays of such interfaces requires much more computational power than a single server has available, and thus single-tier architectures do not scale to thousands of users. The commoditization of the PC and the availability of cheap client computers led to the developlnent of the two-tier architecture.

Two-tier architectures, often also referred to as client-server architectures, consist of a client computer and a server computer, which interact through a well-defined protocol. What part of the functionality the client implements, and what part is left to the server, can vary. In the traditional client-server architecture, the client implements just the graphical user interface, and the server. implements both the business logic and the data management; such clients are often called thin clients, and this architecture is illustra,ted in Figure 7.6.

Other divisions are possible, such as more powerful clients that implement both user interface and business logic, or clients that implement user interface and part of the business logic, with the remaining part being implemented at the

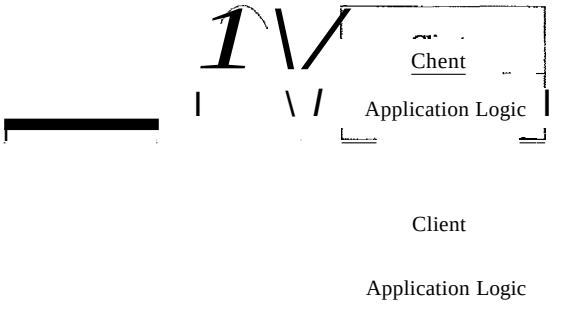


Figure 7.7 A Two-Tier Architecture: Thick Clients

server level; such clients are often called **thick** clients, and this architecture is illustrated in Figure 7.7.

Compared to the single-tier architecture, two-tier architectures physically separate the user interface from the data management layer. To implement two-tier architectures, we can no longer have dumb terminals on the client side; we require computers that run sophisticated presentation code (and possibly, application logic).

Over the last ten years, a large number of client-server development tools such Microsoft Visual Basic and Sybase Powerbuilder have been developed. These tools permit rapid development of client-server software, contributing to the success of the client-server model, especially the thin-client version.

The thick-client model has several disadvantages when compared to the thin-client model. First, there is no central place to update and maintain the business logic, since the application code runs at many client sites. Second, a large amount of trust is required between the server and the clients. As an exam-- pIe, the DBMS of a bank has to trust the (application executing at an) ATM machine to leave the database in a consistent state. (One way to address this problem is through *stored procedures*, trusted application code that is registered with the DBMS and can be called from SQL statelnents. We discuss stored procedures in detail in Section 6.5.)

A third disadvantage of the thick-client architecture is that it does not scale with the number of clients; it typically cannot handle more than a few hundred clients. The application logic at the client issues SQL queries to the server and the server returns the query result to the client, where further processing takes place. Large query results might be transferred between client and server.

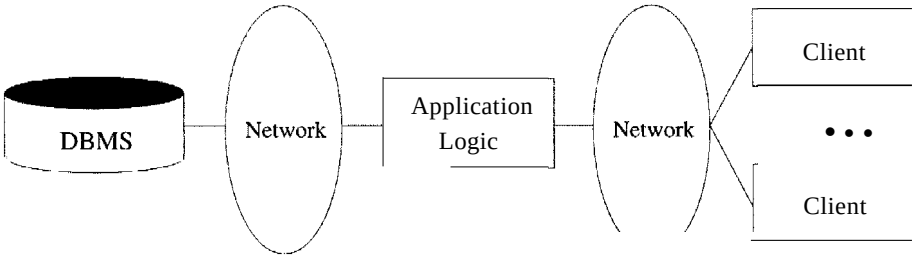


Figure 7.8 A Standard Three-Tier Architecture

(Stored procedures can mitigate this bottleneck.) Fourth, thick-client systems do not scale as the application accesses more and more database systems. Assume there are x different database systems that are accessed by y clients, then there are $x \cdot y$ different connections open at any time, clearly not a scalable solution.

These disadvantages of thick-client systems and the widespread adoption of standard, very thin clients—notably, Web browsers—have led to the widespread use thin-client architectures.

7.5.2 Three-Tier Architectures

The thin-client two-tier architecture essentially separates presentation issues from the rest of the application. The three-tier architecture goes one step further, and also separates application logic from data management:

- Presentation Tier: Users require a natural interface to make requests, provide input, and to see results. The widespread use of the Internet has made Web-based interfaces increasingly popular.
- Middle Tier: The application logic executes here. An enterprise-class application reflects complex business processes, and is coded in a general purpose language such as C++ or Java.
- Data Management Tier: Data-intensive Web applications involve DBMSs, which are the subject of this book.

Figure 7.8 shows a basic three-tier architecture. Different technologies have been developed to enable distribution of the three tiers of an application across multiple hardware platforms and different physical sites. Figure 7.9 shows the technologies relevant to each tier.

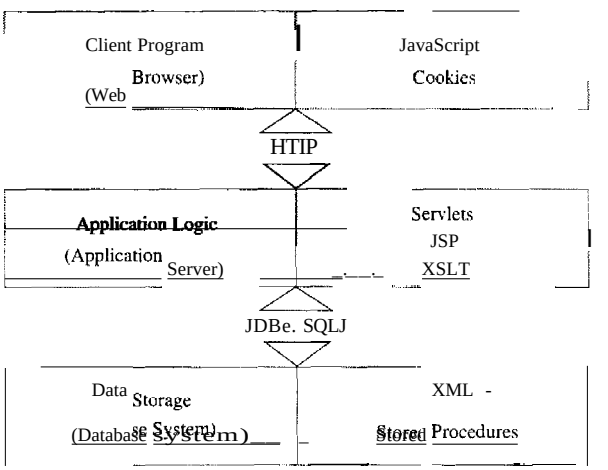


Figure 7.9 Technologies for the Three Tiers

Overview of the Presentation Tier

At the presentation layer, we need to provide forms through which the user can issue requests, and display responses that the middle tier generates. The hypertext markup language (HTML) discussed in Section 7.3 is the basic data presentation language.

It is important that this layer of code be easy to adapt to different display devices and formats; for example, regular desktops versus handheld devices versus cell phones. This adaptivity can be achieved either at the middle tier through generation of different pages for different types of client, or directly at the client through style sheets that specify how the data should be presented. In the latter case, the middle tier is responsible for producing the appropriate data in response to user requests, whereas the presentation layer decides *how* to display that information.

We cover presentation tier technologies, including style sheets, in Section 7.6.

Overview of the Middle Tier

The middle layer runs code that implements the business logic of the application: It controls what data needs to be input before an action can be executed, determines the control flow between multi-action steps, controls access to the database layer, and often assembles dynamically generated HTML pages from database query results.

The middle tier code is responsible for supporting all the different roles involved in the application. For example, in an Internet shopping site implementation, we would like customers to be able to browse the catalog and make purchases, administrators to be able to inspect current inventory, and possibly data analysts to ask summary queries about purchase histories. Each of these roles can require support for several complex actions.

For example, consider the a customer who wants to buy an item (after browsing or searching the site to find it). Before a sale can happen, the customer has to go through a series of steps: She has to add items to her shopping basket, she has to provide her shipping address and credit card number (unless she has an account at the site), and she has to finally confirm the sale with tax and shipping costs added. Controlling the flow among these steps and remembering already executed steps is done at the middle tier of the application. The data carried along during this series of steps might involve database accesses, but usually it is not yet permanent (for example, a shopping basket is not stored in the database until the sale is confirmed).

We cover the middle tier in detail in Section 7.7.

7.5.3 Advantages of the Three-Tier Architecture

The three-tier architecture has the following advantages:

- Heterogeneous Systems: Applications can utilize the strengths of different platforms and different software components at the different tiers. It is easy to modify or replace the code at any tier without affecting the other tiers.
- **Thin** Clients: Clients only need enough computation power for the presentation layer. Typically, clients are Web browsers.
- Integrated Data Access: In many applications, the data must be accessed from several sources. This can be handled transparently at the middle tier, where we can centrally manage connections to all database systems involved.
- Scalabilit,y to Many Clients: Each client is lightweight and all access to the system is through the middle tier. The middle tier can share database connections across clients, and if the middle tier becomes the bottle-neck, we can deploy several servers executing the middle tier code; clients can connect to anyone of these servers, if the logic is designed appropriately. This is illustrated in Figure 7.10, which also shows how the middle tier accesses multiple data sources. Of course, we rely upon the DBMS for each

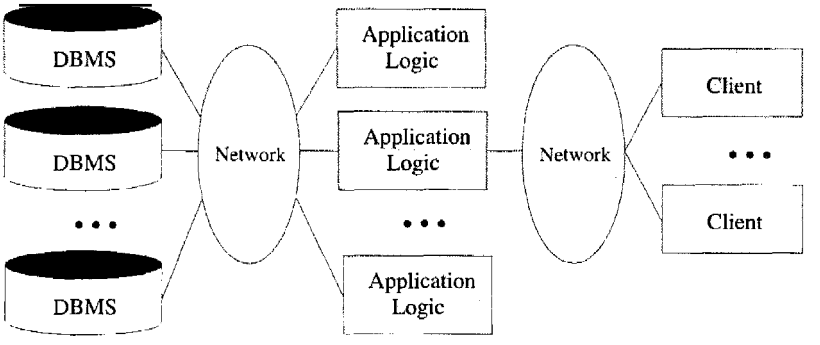


Figure 7.10 Middle-Tier Replication and Access to Multiple Data Sources

data source to be scalable (and this might involve additional parallelization or replication, as discussed in Chapter 22).

- Software Development Benefits: By dividing the application cleanly into parts that address presentation, data access, and business logic, we gain many advantages. The business logic is centralized, and is therefore easy to maintain, debug, and change. Interaction between tiers occurs through well-defined, standardized APIs. Therefore, each application tier can be built out of reusable components that can be individually developed, debugged, and tested.

7.6 THE PRESENTATION LAYER

In this section, we describe technologies for the client side of the three-tier architecture. We discuss HTML forms as a special means of passing arguments from the client to the middle tier (i.e., from the presentation tier to the middle tier) in Section 7.6.1. In Section 7.6.2, we introduce JavaScript, a Java-based scripting language that can be used for light-weight computation in the client tier (e.g., for simple animations). We conclude our discussion of client-side technologies by presenting style sheets in Section 7.6.3. Style sheets are languages that allow us to present the same webpage with different formatting for clients with different presentation capabilities; for example, Web browsers versus cell phones, or even a Netscape browser versus Microsoft's Internet Explorer.

7.6.1 HTML Forms

HTML forms are a common way of communicating data from the client tier to the middle tier. The general format of a form is the following:

```
<FORM ACTION="page.jsp" METHOD="GET" NAME="LoginForm">
```

```
</FORM>
```

A single HTML document can contain more than one form. Inside an HTML form, we can have any HTML tags except another FORM element.

The FORM tag has three important attributes:

- ACTION: Specifies the URI of the page to which the form contents are submitted; if the ACTION attribute is absent, then the URI of the current page is used. In the sample above, the form input would be submited to the page named page. j sp, which should provide logic for processing the input from the form. (We will explain methods for reading form data at the middle tier in Section 7.7.)
- METHOD: The HTTP/1.0 method used to submit the user input from the filled-out form to the webserver. There are two choices, GET and POST; we postpone their discussion to the next section.
- NAME: This attribute gives the form a name. Although not necessary, naming forms is good style. In Section 7.6.2, we discuss how to write client-side programs in JavaScript that refer to forms by name and perform checks on form fields.

Inside HTML forms, the INPUT, SELECT, and TEXTAREA tags are used to specify user input elements; a form can have many elements of each type. The simplest user input element is an INPUT field, a standalone tag with no terminating tag. An example of an INPUT tag is the following:

```
<INPUT TYPE=ltext" NAME="title">
```

The INPUT tag has several attributes. The three most important ones are TYPE, NAME, and VALUE. The TYPE attribute determines the type of the input field. If the TYPE attribute has value **text**, then the field is a text input field. If the TYPE attribute has value password, then the input field is a text field where the entered characters are displayed as stars on the screen. If the TYPE attribute has value **reset**, it is a simple button that resets all input fields within the form to their default values. If the TYPE attribute has value submit, then it is a button that sends the values of the different input fields in the form to the server. Note that **reset** and submit input fields affect the entire form.

The NAME attribute of the INPUT tag specifies the symbolic name for this field and is used to identify the value of this input fi.eld when it is sent to the server. NAME has to be set for INPUT tags of all types except submit and **reset**. In the preceding example, we specified **title** as the NAME of the input field.

The VALUE attribute of an input tag can be used for text or password fields to specify the default contents of the field. For submit or reset buttons, VALUE determines the label of the button.

The form in Figure 7.11 shows two text fields, one regular text input field and one password field. It also contains two buttons, a reset button labeled 'Reset Values' and a submit button labeled 'Log on.' Note that the two input fields are named, whereas the reset and submit button have no NAME attributes.

```
<FORM ACTION="page.jsp" METHoD="GET" NAME="LoginForm">
    <INPUT TYPE="text" NAME="username" VALUE=" Joe"><P>
    <INPUT TYPE="password" NAME="p&ssword"><P>
    <INPUT TYPE="reset" VALUE="Reset Values"><P>
    <INPUT TYPE="submit" VALUE="Log on">
</FoRM>
```

Figure 7.11 HTl'vlL Form with Two Text Fields and Two Buttons

HTML forms have other ways of specifying user input, such as the aforementioned TEXTAREA and SELECT tags; we do not discuss them.

Passing Arguments to Server-Side Scripts

As mentioned at the beginning of Section 7.6.1, there are two different ways to submit HTML Form data to the webserver. If the method GET is used, then the contents of the form are assembled into a query URI (as discussed next) and sent to the server. If the method POST is used, then the contents of the form are encoded as in the GET method, but the contents are sent in a separate data block instead of appending them directly to the URI. Thus, in the GET method the form contents are directly visible to the user as the constructed URI, whereas in the POST method, the form contents are sent inside the HTTP request message body and are not visible to the user.

Using the GET method gives users the opportunity to bookmark the page with the constructed URI and thus directly jump to it in subsequent sessions; this is not possible with the POST method. The choice of GET versus POST should be determined' by the application and its requirements.

Let us look at the encoding of the URI when the GET method is used. The encoded URI has the following form:

action?name1=value1&name2=value2&name3=value3

The action is the URI specified in the ACTION attribute to the FORM tag, or the current document URI if no ACTION attribute was specified. The 'name=value' pairs are the user inputs from the INPUT fields in the form. For form INPUT fields where the user did not input anything, the name is stil present with an empty value (name=). As a concrete example, consider the password submission form at the end of the previous section. Assume that the user inputs 'John Doe' as username, and 'secret' as password. Then the request URI is:

page.jsp?username=J01111+Doe&password=secret

The user input from forms can contain general ASCII characters, such as the space character, but URIs have to be single, consecutive strings with no spaces. Therefore, special characters such as spaces, '=', and other unprintable characters are encoded in a special way. To create a URI that has form fields encoded, we perform the following three steps:

1. Convert all special characters in the names and values to '%xyz,' where 'xyz' is the ASCII value of the character in hexadecimal. Special characters include =, &, %, +, and other unprintable characters. Note that we could encode *all* characters by their ASCII value.
2. Convert all space characters to the '+' character.
3. Glue corresponding names and values from an individual HTML INPUT tag together with '=' and then paste name-value pairs from different HTML INPUT tags together using'&' to create a request URI of the form: action?namel=value1&name2=value2&name3=value3

Note that in order to process the input elements from the HTML form at the middle tier, we need the ACTION attribute of the FORM tag to point to a page, script, or program that will process the values of the form fields the user entered. We discuss ways of receiving values from form fields in Sections 7.7.1 and 7.7.3.

7.6.2 JavaScript

JavaScript is a scripting language at the client tier with which we can add programs to webpages that run directly at the client (Le., at the machine running the Web browser). JavaScript is often used for the following types of computation at the client:

- Browser Detection: JavaScript can be used to detect the browser type and load a browser-specific page.
- Form Validation: JavaScript is used to perform simple consistency checks on form fields. For example, a JavaScript program might check whether a

form input that asks for an email address contains the character '@,' or if all required fields have been input by the user.

- **Browser Control:** This includes opening pages in customized windows; examples include the annoying pop-up advertisements that you see at many websites, which are programmed using JavaScript.

JavaScript is usually embedded into an HTML document with a special tag, the SCRIPT tag. The SCRIPT tag has the attribute LANGUAGE, which indicates the language in which the script is written. For JavaScript, we set the language attribute to JavaScript. Another attribute of the SCRIPT tag is the SRC attribute, which specifies an external file with JavaScript code that is automatically embedded into the HTML document. Usually JavaScript source code files use a '.js' extension. The following fragment shows a JavaScript file included in an HTML document:

```
<SCRIPT LANGUAGE=" JavaScript" SRC="validateForm.js"> </SCRIPT>
```

The SCRIPT tag can be placed inside HTML comments so that the JavaScript code is not displayed verbatim in Web browsers that do not recognize the SCRIPT tag. Here is another JavaScipt code example that creates a pop-up box with a welcoming message. We enclose the JavaScipt code inside HTML comments for the reasons just mentioned.

```
<SCRIPT LANGUAGE=" JavaScript" >
<!--
    alert("Welcome to our bookstore");
//-->
</SCRIPT>
```

JavaScript provides two different commenting styles: single-line comments that start with the '//' character, and multi-line comments starting with '/*' and ending with ,*/' characters.[1]

JavaScript has variables that can be numbers, boolean values (true or false), strings, and some other data types that we do not discuss. Global variables have to be declared in advance of their usage with the keyword var, and they can be used anywhere inside the HTML documents. Variables local to a JavaScript function (explained next) need not be declared. Variables do not have a fixed type, but implicitly have the type of the data to which they have been assigned.

---

1 Actually, '<!--' also marks the start of a single-line comment, which is why we did not have to mark the HTML starting cormnent '<!--' in the preceding example using JavaScript comment notation. In contrast, the HTML closing comment "-->" has to be commented out in JavaScript as it is interpreted otherwise.

JavaScript has the usual assignment operators (=, + =, etc.), the usual arithmetic operators (+, -, *, /, %), the usual comparison operators (==, ! =, >=, etc.), and the usual boolean operators (&& for logical AND, || for logical OR, and! for negation). Strings can be concatenated using the '+' character. The type of an object determines the behavior of operators; for example 1+1 is 2, since we are adding numbers, whereas "1"+"1" is "11," since we are concatenating strings. JavaScript contains the usual types of statements, such as assignments, conditional statements (if Ccondition) {statements;} else {statements; }), and loops (for-loop, do-while, and while-loop).

JavaScript allows us to create functions using the function keyword: function f Cargl, arg2) {statements;}. We can call functions from JavaScript code, and functions can return values using the keyword return.

We conclude this introduction to JavaScript with a larger example of a JavaScript function that tests whether the login and password fields of a HTML form are not empty. Figure 7.12 shows the JavaScript function and the HTML form. The JavaScript code is a function called testLoginEmptyO that tests whether either of the two input fields in the form named LoginForm is empty. In the function testLoginEmpty, we first use variable loginForm to refer to the form LoginForm using the implicitly defined variable document, which refers to the current HTML page. (JavaScript has a library of objects that are implicitly defined.) We then check whether either of the strings loginForm. userif. value or loginForm. password. value is empty.

The function testLoginEmpty is checked within a form event handler. An event **handler** is a function that is called if an event happens on an object in a webpage. The event handler we use is onSubmit, which is called if the submit button is pressed (or if the user presses return in a text field in the form). If the event handler returns true, then the form contents are submitted to the server, otherwise the form contents are not submitted to the server.

JavaScript has functionality that goes beyond the basics that we explained in this section; the interested reader is referred to the bibliographic notes at the end of this chapter.

7.6.3 Style Sheets

Different clients have different displays, and we need correspondingly different ways of displaying the same information. For example, in the simplest case, we might need to use different font sizes or colors that provide high-contrast on a black-and-white screen. As a more sophisticated example, we might need to re-arrange objects on the page to accommodate small screens in personal

```
<SCRIPT LANGUAGE="JavaScript">
<!--
function testLoginEmpty()
{
    10ginForm = document.LoginForm
    if ((loginForm.userid.value == "") ||
            (loginFonn.password.value == "")) {
        alert(,Please enter values for userid and password.');
        return false;
    }
    else
        return true;
}
//-->
</SCRIPT>
<Hi ALIGN = "CENTER">Barns and Nobble Internet Bookstore</Hi>
<H3 ALIGN = "CENTER">Please enter your userid and password:</H3>
<FORM NAME = "LoginForm" METHOD="POST"
        ACTION="TableOfContents.jsp"
        onSubmit="return testLoginEmptyO">
    Userid: <INPUT TYPE="TEXT" NAME="userid"><P>
    Password: <INPUT TYPE="PASSWORD" NAME="password"><P>
    <INPUT TYPE="SUBMIT" VALUE="Login" NAME="SUBMIT">
    <INPUT TYPE="RESET" VALUE="Clear Input" NAME="RESET">
</FORM>
```

Figure 7.12 Form Validation with JavaScript

digital assistants (PDAs). As another example, we might highlight different infonnation to focus on some important part of the page. A style sheet is a method to adapt the same document contents to different presentation formats. A style sheet contains instructions that tell a Web browser (or whatever the client uses to display the webpage) how to translate the data of a document into a presentation that is suitable for the client's display.

Style sheets separate the transformative aspect of the page from the rendering aspects of the page. During transformation, the objects in the XML document are rearranged to form a different structure, to omit parts of the XML document, or to merge two different XML documents into a single document. During rendering, we take the existing hierarchical structure of the XML document and format the document according to the user's display device.

```
BODY {BACKGROUND-COLOR: yellow}
Hi {FONT-SIZE: 36pt}
H3 {COLOR: blue}
P {MARGIN-LEFT: 50px; COLOR: red}
```

Figure 7.13 An Example Style sheet

The use of style sheets has many advantages. First, we can reuse the same document many times and display it differently depending on the context. Second, we can tailor the display to the reader's preference such as font size, color style, and even level of detail. Third, we can deal with different output formats, such as different output devices (laptops versus cell phones), different display sizes (letter versus legal paper), and different display media (paper versus digital display). Fourth, we can standardize the display format within a corporation and thus apply style sheet conventions to documents at any time. Further, changes and improvements to these display conventions can be managed at a central place.

There are two style sheet languages: XSL and ess. ess was created for HTML with the goal of separating the display characteristics of different formatting tags from the tags themselves. XSL is an extension of ess to arbitrary XML docurnents; besides allowing us to define ways of formatting objects, XSL contains a transformation language that enables us to rearrange objects. The target files for ess are HTML files, whereas the target files for XSL are XML files.

Cascading Style Sheets

A Cascading Style Sheet (CSS) defines how to display HTML elements. (In Section 7.13, we introduce a more general style sheet language designed for XML documents.) Styles are normally stored in style sheets, which are files that contain style definitions. Many different HTML documents, such as all documents in a website, can refer to the same ess. Thus, we can change the format of a website by changing a single file. This is a very convenient way of changing the layout of many webpages at the same time, and a first step toward the separation of content from presentation.

An example style sheet is shown in Figure 7.13. It is included into an HTML file with the following line:

```
<LINK REL="style sheet" TYPE="text/css" HREF="books.css" />
```

Each line in a CSS sheet consists of three parts; a selector, a property, and a value. They are syntactically arranged in the following way:

selector {property: value}

The selector is the element or tag whose format we are defining. The property indicates the tag's attribute whose value we want to set in the style sheet, and the property is the actual value of the attribute. As an example, consider the first line of the example style sheet shown in Figure 7.13:

BODY {BACKGROUND-COLOR: yellow}

This line has the same effect as changing the HTML code to the following:

<BODY BACKGROUND-COLOR="yellow">.

The value should always be quoted, as it could consist of several words. More than one property for the same selector can be separated by semicolons as shown in the last line of the example in Figure 7.13:

P {MARGIN-LEFT: 50px; COLOR: red}

Cascading style sheets have an extensive syntax; the bibliographic notes at the end of the chapter point to books and online resources on CSSs.

XSL

XSL is a language for expressing style sheets. An XSL style sheet is, like CSS, a file that describes how to display an XML document of a given type. XSL shares the functionality of CSS and is compatible with it (although it uses a different syntax).

The capabilities of XSL vastly exceed the functionality of CSS. XSL contains the XSL Transformation language, or XSLT, a language that allows us to transform the input XML document into a XML document with another structure. For example, with XSLT we can change the order of elements that we are displaying (e.g.; by sorting them), process elements more than once, suppress elements in one place and present them in another, and add generated text to the presentation.

XSL also contains the XML Path Language (XPath), a language that allows us to refer to parts of an XML document. We discuss XPath in Section

27. XSL also contains XSL Formatting Object, a way of formatting the output of an XSL transformation.

7.7 THE MIDDLE TIER

In this section, we discuss technologies for the middle tier. The first generation of middle-tier applications were stand-alone programs written in a general-purpose programming language such as C, C++, and Perl. Programmers quickly realized that interaction with a stand-alone application was quite costly; the overheads include starting the application every time it is invoked and switching processes between the webserver and the application. Therefore, such interactions do not scale to large numbers of concurrent users. This led to the development of the application server, which provides the run-time environment for several technologies that can be used to program middle-tier application components. Most of today's large-scale websites use an application server to run application code at the middle tier.

Our coverage of technologies for the middle tier mirrors this evolution. We start in Section 7.7.1 with the Common Gateway Interface, a protocol that is used to transmit arguments from HTML forms to application programs running at the middle tier. We introduce application servers in Section 7.7.2. We then describe technologies for writing application logic at the middle tier: Java servlets (Section 7.7.3) and Java Server Pages (Section 7.7.4). Another important functionality is the maintenance of state in the middle tier component of the application as the client component goes through a series of steps to complete a transaction (for example, the purchase of a market basket of items or the reservation of a flight). In Section 7.7.5, we discuss Cookies, one approach to maintaining state.

7.7.1 CGI: The Common Gateway Interface

The Common Gateway Interface connects HTML forms with application programs. It is a protocol that defines how arguments from forms are passed to programs at the server side. We do not go into the details of the actual CGI protocol since libraries enable application programs to get arguments from the HTML fonn; we shortly see an example in a CGI program. Programs that communicate with the webserver via CGI are often called CGI scripts, since many such application programs were written in a scripting language such as Perl.

As an example of a program that interfaces with an HTML form via CGI, consider the sample page shown in Figure 7.14. This webpage contains a form where a user can fill in the name of an author. If the user presses the 'Send

```
<HTML><HEAD><TITLE>The Database Bookstore</TITLE></HEAD>
<BODY>
<FORM ACTION="find_books.cgi" METHOD=POST>
    Type an author name:
    <INPUT TYPE="text" NAME=lauthorName"
           SIZE=30 MAXLENGTH=50>
    <INPUT TYPE="submitil value="Send it">
    <INPUT TYPE=lreset" VALUE="Clear form">
</FORM>
</BODY></HTML>
```

Figure 7.14 A Sample Web Page Where Form Input Is Sent to a CGI Script

it' button, the Perl script 'findBooks.cgi' shown in Figure 7.14 is executed as a separate process. The CGl protocol defines how the communication between the form and the script is performed. Figure 7.15 illustrates the processes created when using the CGl protocol.

Figure 7.16 shows the example CGl script, written in Perl. We omit error-checking code for simplicity. Perl is an interpreted language that is often used for CGl scripting and many Perl libraries, called **modules,** provide high-level interfaces to the CGl protocol. \Ve use one such library, called the **DBI library,** in our example. The CGI module is a convenient collection of functions for creating CGl scripts. In part 1 of the sample script, we extract the argument of the HTML form that is passed along from the client as follows:

```
$authorName = $dataIn- >paramCauthorName');
```

Note that the parameter name authorName was used in the form in Figure 7.14 to name the first input field. Conveniently, the CGl protocol abstracts the actual implementation of how the webpage is returned to the Web browser; the webpage consists simply of the output of our program, and we start assembling the output HTML page in part 2. Everything the script writes in print-statements is part of the dynamically constructed webpage returned to the browser. We finish in part 3 by appending the closing format tags to the resulting page.

7.7.2 Application Servers

Application logic can be enforced through server-side programs that are invoked using the CGl protocol. However, since each page request results in the creation of a new process, this solution does not scale well to a large number of simultaneous requests. This performance problem led to the development of

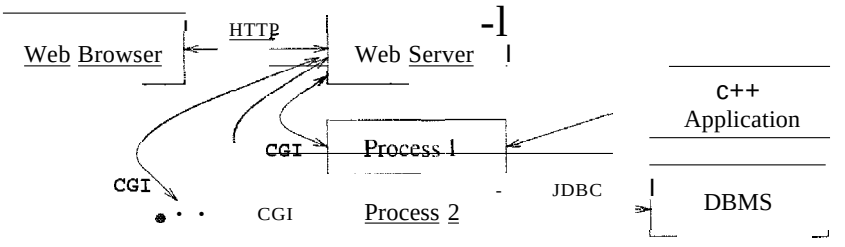


Figure 7.15 Process Structure with eGI Scripts

```
#!/usr/bin/perl
use CGI;

### part 1
$dataln = new CGI;
$dataIn-¿header();
$authorName = $dataln-l,param('authorName');

### part 2
print (II<HTML><TITLE>Argument passing test</TITLE> II) ;
print (II The user passed the following argument: II) ;
print (II authorName: ", $authorName);

### part 3
print ("</HTML>");
exit;
```

Figure 7.16 A Simple Perl Script

specialized programs called application servers. An application server maintains a pool of threads or processes and uses these to execute requests. Thus, it avoids the startup cost of creating a new process for each request.

Application servers have evolved into flexible middle-tier packages that provide many functions in addition to eliminating the process-creation overhead. They facilitate concurrent access to several heterogeneous data sources (e.g., by providing JDBC drivers), and provide session management services. Often, business processes involve several steps. Users expect the system to maintain continuity during such a multistep session. Several session identifiers such as cookies, URI extensions, and hidden fields in HTML forms can be used to identify a session. Application servers provide functionality to detect when a session starts and ends and keep track of the sessions of individual users. They

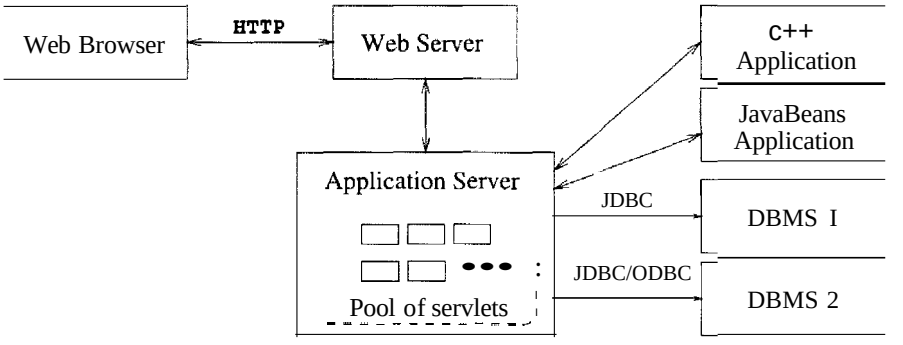


Figure 7.17 Process Structure in the Application Server Architecture

also help to ensure secure database access by supporting a general user-id mechanism. (For more on security, see Chapter 21.)

A possible architecture for a website with an application server is shown in Figure 7.17. The client (a Web browser) interacts with the webserver through the HTTP protocol. The webserver delivers static HTML or XML pages directly to the client. To assemble dynamic pages, the webserver sends a request to the application server. The application server contacts one or more data sources to retrieve necessary data or sends update requests to the data sources. After the interaction with the data sources is completed, the application server assembles the webpage and reports the result to the webserver, which retrieves the page and delivers it to the client.

The execution of business logic at the webserver's site, server-side processing, has become a standard model for implementing more complicated business processes on the Internet. There are many different technologies for server-side processing and we only mention a few in this section; the interested reader is referred to the bibliographic notes at the end of the chapter.

7.7.3 Servlets

Java servlets are pieces of Java code that run on the middle tier, in either webservers or application servers. There are special conventions on how to read the input from the user request and how to write output generated by the servlet. Servlets are truly platform-independent, and so they have become very popular with Web developers.

Since servlets are Java programs, they are very versatile. For example, servlets can build webpages, access databases, and maintain state. Servlets have access

```
import java.io.*;
import javax.servlet.*;
import javax.servlet.http.*;

pUblic class ServletTemplate extends HttpServlet {
    public void doGet(HttpServletRequest request,
            HttpServletResponse response)
            throws ServletException, IOException {
        PrintWriter out = response.getWriter();
        // Use 'out' to send content to browser
        out.println("Hello World");
    }
}
```

Figure 7.18 Servlet Template

to all Java APIs, including JDBC. All servlets must implement the Servlet interface. In most cases, servlets extend the specific HttpServlet class for servers that communicate with clients via HTTP. The HttpServlet class provides methods such as doGet and doPost to receive arguments from HTML forms, and it sends its output back to the elient via HTTP. Servlets that communicate through other protocols (such as ftp) need to extend the class GenericServlet.

Servlets are compiled Java classes executed and maintained by a servlet container. The servlet container manages the lifespan of individual servlets by creating and destroying them. Although servlets can respond to any type of request, they are commonly used to extend the applications hosted by webservers. For such applications, there is a useful library of HTTP-specific servlet classes.

Servlets usually handle requests from HTML forms and maintain state between the client and the server. We discuss how to maintain state in Section 7.7.5. A template of a generic servlet structure is shown in Figure 7.18. This simple servlet just outputs the two words "Hello World," but it shows the general structure of a full-fledged servlet. The request object is used to read HTML form data. The response object is used to specify the HTTP response status code and headers of the HTTP response. The object out is used to compose the content that is returned to the client.

Recall that HTTP sends back the status line, a header, a blank line, and then the context. Right now our servlet just returns plain text. We can extend our servlet by setting the content type to HTML, generating HTML as follows:

```
PrinfWriter out = response.get\Vriter();
String docType =
    "<!DOCTYPE HTML PUBLIC "-//W3C//DTD HTML 4.0 " +
    "Transitional//EN"> \n";
out.println(docType +
    "<HTML>\n" +
    "<HEAD><TITLE>Hello WWW</TITLE></HEAD>\n" +
    "<BODY>\n" +
    "<Hl>Hello WWW</Hl>\n" +
    "</BODY></HTML>");
```

What happens during the life of a servlet? Several methods are called at different stages in the development of a servlet. When a requested page is a servlet, the webserver forwards the request to the servlet container, which creates an instance of the servlet if necessary. At servlet creation time, the servlet container calls the init () method, and before deallocating the servlet, the servlet container calls the servlet's destroyO method.

When a servlet container calls a servlet because of a requested page, it starts with the service () method, whose default behavior is to call one of the following methods based on the HTTP transfer method: service () calls doGet 0 for a HTTP GET request, and it calls doPost () for a HTTP POST request. This automatic dispatching allows the servlet to perform different tasks on the request data depending on the HTTP transfer method. Usually, we do not override the service () method, unless we want to program a servlet that handles both HTTP POST and HTTP GET requests identically.

We conclude our discussion of servlets with an example, shown in Figure 7.19, that illustrates how to pass arguments from an HTML form to a servlet.

7.7.4 JavaServer Pages

In the previous section, we saw how to use Java programs in the middle tier to encode application logic and dynamically generate webpages. If we needed to generate HTML output, we wrote it to the out object. Thus, we can think about servlets as Java code embodying application logic, with embedded HTML for output.

JavaServer pages (JSPs) interchange the roles of output and application logic. JavaServer pages are written in HTML with servlet-like code embedded in special HT1VIL tags. Thus, in comparison to servlets, JavaServer pages are better suited to quickly building interfaces that have some logic inside, whereas servlets are better suited for complex application logic.

```
import java.io.*;
import javax.servlet.*;
import javax.servlet.http.*;
import java.util.*;

public class ReadUserName extends HttpServlet {
    public void doGet(HttpServletRequest request,
            HttpServletResponse response)
        throws ServletException, IOException {

        response.setContentType('j textjhtml'j);
        PrintWriter out = response.getWriter();

        out.println("<BODY>\n" +
            "<Hi ALIGN=CENTER> Username: </Hi>\n" +
            "<UL>\n" +
            " <LI>title: "
            + request.getParameter("userid") + "\n" +
            + request.getParameter("password'j) + "\n" +
            1</UL>\n" +
            1</BODY></HTML>")j
    }

    public void doPost(HttpServletRequest request,
            HttpServletResponse response)
        throws ServletException, IOException {
        doGet(request, response);
    }
}
```

Figure 7.19 Extracting the User Name and Password From a Form

While there is a big difference for the programmer, the middle tier handles JavaServer pages in a very simple way: They are usually compiled into a servlet, which is then handled by a servlet container analogous to other servlets.

The code fragment in Figure 7.20 shows a simple JSP example. In the middle of the HTML code, we access information that was passed from a form.

```
<!DOCTYPE HTML PUBLIC 11_//W3C//DTD HTML 4.0
        Transitional//EN"">
<HTML>
<HEAD><TITLE>Welcome to Barnes and Nobble</TITLE></HEAD>
<BODY>
    <Hl>Welcome back!</Hl>
    <% String name="NewUser"";
        if (request.getParameter(lIusernamell) != null) {
            name=request.getParameter("username");
        }
    %>
    You are logged on as user <%=name%>
    <P>
    Regular HTML for all the rest of the on-line store's webpage.
</BODY>
</HTML>
```

Figure 7.20 Reading Form Parameters in JSP

7.7.5 **Maintaining State**

As discussed in previous sections, there is a need to maintain a user's state across different pages. As an example, consider a user who wants to make a purchase at the Barnes and Nobble website. The user must first add items into her shopping basket, which persists while she navigates through the site. Thus, we use the notion of state mainly to remember information as the user navigates through the site.

The HTTP protocol is stateless. We call an interaction with a webserver stateless if no inforination is retained from one request to the next request. We call an interaction with a webserver stateful, or we say that state is maintained, if some memory is stored between requests to the server, and different actions are taken depending on the contents stored.

In our example of Barnes and Nobble, we need to maintain the shopping basket of a user. Since state is not encapsulated in the HTTP protocol, it has to be maintained either at the server or at the client. Since the HTTP protocol is stateless by design, let us review the advantages and disadvantages of this design decision. First, a stateless protocol is easy to program and use, and it is great for applications that require just retrieval of static information. In addition, no extra memory is used to maintain state, and thus the protocol itself is very efficient. On the other hand, without some additional mechanism at the presentation tier and the middle tier, we have no record of previous requests, and we cannot program shopping baskets or user logins.

Since we cannot maintain state in the HTTP protocol, where should we mtain-tain state? There are basically two choices. We can maintain state in the middle tier, by storing information in the local main memory of the application logic, or even in a database system. Alternatively, we can maintain state on the client side by storing data in the form of a *cookie*. We discuss these two ways of maintaining state in the next two sections.

Maintaining State at the Middle Tier

At the middle tier, we have several choices as to *where* we maintain state. First, we could store the state at the bottom tier, in the database server. The state survives crashes of the system, but a database access is required to query or update the state, a potential performance bottleneck. An alternative is to store state in main memory at the middle tier. The drawbacks are that this information is volatile and that it might take up a lot of main memory. We can also store state in local files at the middle tier, as a compromise between the first two approaches.

A rule of thumb is to use state maintenance at the middle tier or database tier only for data that needs to persist over many different user sessions. Examples of such data are past customer orders, click-stream data recording a user's movement through the website, or other permanent choices that a user makes, such as decisions about personalized site layout, types of messages the user is willing to receive, and so on. As these examples illustrate, state information is often centered around users who interact with the website.

Maintaining State at the Presentation Tier: Cookies

Another possibility is to store state at the presentation tier and pass it to the middle tier with every HTTP request. We essentially work around around the statelessness of the HTTP protocol by sending additional information with every request. Such information is called a cookie.

A **cookie** is a collection of *(name, value)*-pairs that can be manipulated at the presentation and middle tiers. Cookies are easy to use in Java servlets and Java8erver Pages and provide a simple way to make non-essential data persistent at the client. They survive several client sessions because they persist in the browser cache even after the browser is closed.

One disadvantage of cookies is that they are often perceived as as being invasive, and many users disable cookies in their Web browser; browsers allow users to prevent cookies from being saved on their machines. Another disadvantage is that the data in a cookie is currently limited to 4KB, but for most applications this is not a bad limit.

We can use cookies to store information such as the user's shopping basket, login information, and other non-permanent choices made in the current session.

Next, we discuss how cookies can be manipulated from servlets at the middle tier.

The Servlet Cookie API

A cookie is stored. in a small text file at the client and. contains *(name, value)*-pairs, where both name and value are strings. We create a new cookie through the Java Cookie class in the middle tier application code:

```
Cookie cookie = new Cookie("username","guest");
cookie.setDomain("www.bookstore.com");
cookie.set8ecure(false);                    // no 88L required
cookie.setMaxAge(60*60*24*7*31);      // one month lifetime
response.addCookie(cookie);
```

Let us look at each part of this code. First, we create a new Cookie object with the specified *(name, value)*-pair. Then we set attributes of the cookie; we list some of the most common attributes below:

- setDomain and getDomain: The domain specifies the website that will receive the cookie. The default value for this attribute is the domain that created the cookie.
- setSecure and getSecure: If this flag is true, then the cookie is sent only if we are llsing a secure version of the HTTP protocol, such as 88L.
- setMaxAge and getMaxAge: The MaxAge attribute determines the lifetime of the cookie in seconds. If the value of MaxAge is less than or equal to zero, the cookie is deleted when the browser is closed.

- setName and getName: We did not use these functions in our code fragment; they allow us to Ilame the cookie.
- setValue and getValue: These functions allow us to set and read the value of the cookie.

The cookie is added to the request object within the Java servlet to be sent to the client. Once a cookie is received from a site (www.bookstore.comin this example), the client's Web browser appends it to all HTTP requests it sends to this site, until the cookie expires.

We can access the contents of a cookie in the middle-tier code through the request object getCookies 0 method, which returns an array of Cookie objects. The following code fragment reads the array and looks for the cookie with name 'username.'

```
Cookie[] cookies = request.getCookiesO;
String theUser;
for(int i=O; i < cookies.length; i++) {
    Cookie cookie = cookies[i];
    if (cookie.getNameO.equals("username"))
        theUser = cookie.getValueO;
}
```

A simple test can be used to check whether the user has turned off cookies: Send a cookie to the user, and then check whether the request object that is returned still contains the cookie. Note that a cookie should never contain an unencrypted password or other private, unencrypted data, as the user can easily inspect, modify, and erase any cookie at any time, including in the middle of a session. The application logic needs to have sufficient consistency checks to ensure that the data in the cookie is valid.

7.8 CASE STUDY: THE INTERNET BOOK SHOP

DBDudes now moves on to the implementation of the application layer and considers alternatives for connecting the DBMS to the World Wide Web.

DBDudes begifls by considering session management. For example, users who log in to the site, browse the catalog, and select books to buy do not want to re-enter their cllstomer identification numbers. Session management has to extend to the whole process of selecting books, adding them to a shopping cart, possibly removing books from the cart, and checking out and paying for the books.

DBDudes then considers whether webpages for books should be static or dynamic. If there is a static webpage for each book, then we need an extra database field in the Books relation that points to the location of the file. Even though this enables special page designs for different books, it is a very labor-intensive solution. DBDudes convinces B&N to dynamically assemble the webpage for a book from a standard template instantiated with information about the book in the Books relation. Thus, DBDudes do not use static HTML pages, such as the one shown in Figure 7.1, to display the inventory.

DBDudes considers the use of XML as a data exchange format between the database server and the middle tier, or the middle tier and the client tier. Representation of the data in XML at the middle tier as shown in Figures 7.2 and 7.3 would allow easier integration of other data sources in the future, but B&N decides that they do not anticipate a need for such integration, and so DBDudes decide not to use XML data exchange at this time.

DBDudes designs the application logic as follows. They think that there will be four different webpages:

- index.jsp: The home page of Barns and Nobble. This is the main entry point for the shop. This page has search text fields and buttons that allow the user to search by author name, ISBN, or title of the book. There is also a link to the page that shows the shopping cart, cart.jsp.
- login.jsp: Allows registered users to log in. Here DBDudes use an HTML form similar to the one displayed in Figure 7.11. At the middle tier, they use a code fragment similar to the piece shown in Figure 7.19 and JavaServerPages as shown in Figure 7.20.
- search.jsp: Lists all books in the database that match the search condition specified by the user. The user can add listed items to the shopping basket; each book has a button next to it that adds it. (If the item is already in the shopping basket, it increments the quantity by one.) There is also a counter that shows the total number of items currently in the shopping basket. (DBDucles makes a note that that a quantity of five for a single item in the shopping basket should indicate a total purchase quantity of five as well.) The search.jsp page also contains a button that directs the user to cart.jsp.
- cart.jsp: Lists all the books currently in the shopping basket. The listing should include all items in the shopping basket with the product name, price, a text box for the quantity (which the user can use to change quantities of items), and a button to remove the item from the shopping basket. This page has three other buttons: one button to continue shopping (which returns the user to page index.jsp), a second button to update the shop-

ping basket with the altered quantities from the text boxes, and a third button to place the order, which directs the user to the page confirm.jsp.

- coni irm. j sp: Lists the complete order so far and allows the user to enter his or her contact information or customer ID. There are two buttons on this page: one button to cancel the order and a second button to submit the final order. The cancel button ernpties the shopping basket and returns the user to the home page. The submit button updates the database with the new order, empties the shopping basket, and returns the user to the home page.

DBDudes also considers the use of JavaScript at the presentation tier to check user input before it is sent to the middle tier. For example, in the page login. j sp, DBDudes is likely to write JavaScript code similar to that shown in Figure 7.12.

This leaves DBDudes with one final decision: how to connect applications to the DBMS. They consider the two main alternatives presented in Section 7.7: CGI scripts versus using an application server infrastructure. If they use CGI scripts, they would have to encode session management logic-not an easy task. If they use an application server, they can make use of all the functionality that the application server provides. Therefore, they recommend that B&N implement server-side processing using an application server.

B&N accepts the decision to use an application server, but decides that no code should be specific to any particular application server, since B&N does not want to lock itself into one vendor. DBDudes agrees proceeds to build the following pieces:

- DBDudes designs top level pages that allow customers to navigate the website as well as various search forms and result presentations.
- Assuming that DBDudes selects a Java-based application server, they have to write Java servlets to process form-generated requests. Potentially, they could reuse existing (possibly commercially available) JavaBeans. They can use JDBC as a database interface; exarnples of JDBC code can be found in Section 6.2. Instead of prograrnming servlets, they could resort to Java Server Pages and annotate pages with special .JSP markup tags.
- DBDudes select an application server that uses proprietary markup tags, but due to their arrangement with B&N, they are not allowed to use such tags in their code.

For completeness, we remark that if DBDudes and B&N had agreed to use CGr scripts, DBDucles would have had the following tasks:

- Create the top level HTML pages that allow users to navigate the site and various forms that allow users to search the catalog by ISBN, author name, or title. An example page containing a search form is shown in Figure 7.1. In addition to the input forms, DBDudes must develop appropriate presentations for the results.
- Develop the logic to track a customer session. Relevant information must be stored either at the server side or in the customer's browser using cookies.
- Write the scripts that process user requests. For example, a customer can use a form called 'Search books by title' to type in a title and search for books with that title. The CGI interface communicates with a script that processes the request. An example of such a script written in Perl using the DBI library for data access is shown in Figure 7.16.

Our discussion thus far covers only the customer interface, the part of the website that is exposed to B&N's customers. DBDudes also needs to add applications that allow the employees and the shop owner to query and access the database and to generate summary reports of business activities.

Complete files for the case study can be found on the webpage for this book.

7.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are URIs and URLs? (Section 7.2.1)
- How does the HTTP protocol work? What is a stateless protocol? (Section 7.2.2)
- Explain the main concepts of HTML. Why is it used only for data presentation and not data exchange? (Section 7.3)
- What are some shortcomings of HTML, and how does XML address them? (Section 7.4)
- What are the main components of an XML document? (Section 7.4.1)
- Why do we have XML DTDs? What is a well-formed XML document? What is a valid XML document? Give an example of an XML document that is valid but not well-formed, and vice versa. (Section 7.4.2)
- 'What is the role of domain-specific DTDs? (Section 7.4.3)
- What is a three-tier architecture? 'What advantages does it offer over single-tier and two-tier architectures? Give a short overview of the functionality at each of the three tiers. (Section 7.5)

- Explain how three-tier architectures address each of the following issues of database-backed Internet applications: heterogeneity, thin clients, data integration, scalability, software development. (Section 7.5.3)
- Write an HTML form. Describe all the components of an HTML form. (Section 7.6.1)
- What is the difference between the HTML GET and POST methods? How does URI encoding of an HTML form work? (Section 7.11)
- What is JavaScript used for? Write a JavaScipt function that checks whether an HTML form element contains a syntactically valid email address. (Section 7.6.2)
- What problem do style sheets address? What are the advantages of using style sheets? (Section 7.6.3)
- What are Cascading Style Sheets? Explain the components of Cascading Style Sheets. What is XSL and how it is different from CSS? (Sections 7.6.3 and 7.13)
- What is CGl and what problem does it address? (Section 7.7.1)
- What are application servers and how are they different from webservers? (Section 7.7.2)
- What are servlets? How do servlets handle data from HTML forms? Explain what happens during the lifetime of a servlet. (Section 7.7.3)
- What is the difference between servlets and JSP? When should we use servlets and when should we use JSP? (Section 7.7.4)
- Why do we need to maintain state at the middle tier? What are cookies? How does a browser handle cookies? How can we access the data in cookies from servlets? (Section 7.7.5)

EXERCISES

Exercise 7.1 Briefly answer the following questions:

1. Explain the following terms and describe what they are used for: HTML, URL, XML, Java, JSP, XSL, XSLT, servlet, cookie, HTTP, ess, DTD.
2. What is eGl? Why was eGI introduced? What are the disadvantages of an architecture using eel scripts?
3. \Vhat is the difference between a webserver and an application server? What fUl1cionality do typical application servers provide?
4. When is an XML document well-formed? When is an XML document valid?

Exercise 7.2 Briefly answer the following questions about the HTTP protocol:

1. What is a communication protocol?
2. "What is the structure of an HTTP request message? What is the structure of an HTTP response message? Why do HTTP messages carry a version field?
3. What is a stateless protocol? Why was HTTP designed to be stateless?
4. Show the HTTP request message generated when you request the home page of this book (`http://www.cs.wisc.edu/~dbbook`). Show the HTTP response message that the server generates for that page.

Exercise 7.3 In this exercise, you are asked to write the functionality of a generic shopping basket; you will use this in several subsequent project exercises. Write a set of JSP pages that displays a shopping basket of items and allows users to add, remove, and change the quantity of items. To do this, use a cookie storage scheme that stores the following information:

- The UserId of the user who owns the shopping basket.
- The number of products stored in the shopping basket.
- A product id and a quantity for each product.

When manipulating cookies, remember to set the Expires property such that the cookie can persist for a session or indefinitely. Experiment with cookies using JSP and make sure you know how to retrieve, set values, and delete the cookie.

You need to create five JSP pages to make your prototype complete:

- Index Page (index.jsp): This is the main entry point. It has a link that directs the user to the Products page so they can start shopping.
- Products Page (products.jsp): Shows a listing of all products in the database with their descriptions and prices. This is the main page where the user fills out the shopping basket. Each listed product should have a button next to it, which adds it to the shopping basket. (If the item is already in the shopping basket, it increments the quantity by one.) There should also be a counter to show the total number of items currently in the shopping basket. Note that if a user has a quantity of five of a single item in the shopping basket, the counter should indicate a total quantity of five. The page also contains a button that directs the user to the Cart page.
- Cart Page (cart.jsp): Shows a listing of all items in the shopping basket cookie. The listing for each item should include the product name, price, a text box for the quantity (the user can change the quantity of items here), and a button to remove the item from the shopping basket. This page has three other buttons: one button to continue shopping (which returns the user to the Products page), a second button to update the cookie with the altered quantities from the text boxes, and a third button to place or confirm the order, which directs the user to the Confirm page.
- Confirm Page (confirm.jsp): Lists the final order. There are two buttons on this page. One button cancels the order and the other submits the completed order. The cancel button just deletes the cookie and returns the user to the Index page. The submit button updates the database with the new order, deletes the cookie, and returns the user to the Index page.

Exercise 7.4 In the previous exercise, replace the page products.jsp with the following *search page* search.jsp. This page allows users to search products by name or description. There should be both a text box for the search text and radio buttons to allow the

user to choose between search-by-name and search-by-description (as well as a submit button to retrieve the results). The page that handles search results should be modeled after products.jsp (as described in the previous exercise) and be called products.jsp. It should retrieve all records where the search text is a substring of the name or description (as chosen by the user). To integrate this with the previous exercise, simply replace all the links to products.jsp with search.jsp.

Exercise 7.5 'Write a simple authentication mechanism (without using encrypted transfer of passwords, for simplicity). We say a user is authenticated if she has provided a valid username-password combination to the system; otherwise, we say the user is not authenticated. Assume for simplicity that you have a database schema that stores only a customer id and a password:

Passwords(cid: integer, username: string, password: string)

1. How and where are you going to track when a user is 'logged on' to the system?
2. Design a page that allows a registered user to log on to the system.
3. Design a page header that checks whether the user visiting this page is logged in.

Exercise 7.6 (Due to Jeff Derstadt) TechnoBooks.com is in the process of reorganizing its website. A major issue is how to efficiently handle a large number of search results. In a human interaction study, it found that modem users typically like to view 20 search results at a time, and it would like to program this logic into the system. Queries that return batches of sorted results are called *top N queries*. (See Section 25.5 for a discussion of database support for top N queries.) For example, results 1-20 are returned, then results 21-40, then 41-60, and so on. Different techniques are used for performing top N queries and TechnoBooks.com would like you to implement two of them.

Infrastructure: Create a database with a table called Books and populate it with some books, using the format that follows. This gives you 111 books in your database with a title of AAA, BBB, CCC, DDD, or EEE, but the keys are not sequential for books with the same title.

Books(*bookid:* INTEGER, *title:* CHAR(80), *author:* CHAR(80), *price:* REAL)

```
For i = 1 to 111 {
        Insert the tuple (i, "AAA", "AAA Author", 5.99)
        i=i+1
        Insert the tuple (i, "BBB", "BBB Author", 5.99)
        i = i + 1
        Insert the tuple (i, "CCC", "CCC Author", 5.99)
        i=i+1
        Insert the tuple (i, "DDD", "DDD Author", 5.99)
        1=i+1
        Insert the tuple (i, "EEE", "EEE Author", 5.99)
```

Placeholder Technique: The simplest approach to top N queries is to store a placeholder for the first and last result tuples, and then perform the same query. When the new query results are returned, you can iterate to the placeholders and return the previous or next 20 results.

Tuples Shown	Lower Placeholder	Previous Set	Upper Placeholder	Next Set
1-20	1	None	20	21-40
21-40	21	1-20	40	41-60
41-60	41	21-40	60	61-80

Write a webpage in JSP that displays the contents of the Books table, sorted by the Title and BookId, and showing the results 20 at a time. There should be a link (where appropriate) to get the previous 20 results or the next 20 results. To do this, you can encode the placeholders in the Previous or Next Links as follows. Assume that you are displaying records 21-40. Then the previous link is display.jsp?lower=21 and the next link is display.jsp?upper=40.

You should not display a previous link when there are no previous results; nor should you show a Next link if there are no more results. When your page is called again to get another batch of results, you can perform the same query to get all the records, iterate through the result set until you are at the proper starting point, then display 20 more results.

What are the advantages and disadvantages of this technique?

Query Constraints Technique: A second technique for performing top N queries is to push boundary constraints into the query (in the WHERE clause) so that the query returns only results that have not yet been displayed. Although this changes the query, fewer results are returned and it saves the cost of iterating up to the boundary. For example, consider the following table, sorted by (title, primary key).

Batch	Result Number	Title	Primary Key
1	1	AAA	105
1	2	BBB	13
1	3	eee	48
1	4	DDD	52
1	5	DDD	101
2	6	DDD	121
2	7	EEE	19
2	8	EEE	68
2	9	FFF	2
2	10	FFF	33
3	11	FFF	58
3	12	FFF	59
3	13	GGG	93
3	14	EHH	132
3	15	HHH	135

In batch 1, rows 1 through 5 are displayed, in batch 2 rows 6 through 10 are displayed, and so on. Using the placeholder technique, all 15 results would be returned for each batch. Using the constraint technique, batch 1 displays results 1-5 but returns results 1-15, batch 2 will display results 6-10 but returns only results 6-15, and batch 3 will display results 11-15 but return only results 11-15.

The constraint can be pushed into the query because of the sorting of this table. Consider the following query for batch 2 (displaying results 6-10):

```
EXEC SQL SELECT B.Title
FROM     Books B
WHERE    (B.Title = 'DDD' AND B.BookId > 101) OR (B.Title > 'DDD')
ORDER BY B.Title, B.Bookld
```

This query first selects all books with the title 'DDD,' but with a primary key that is greater than that of record 5 (record 5 has a primary key of 101). This returns record 6. Also, any book that has a title after 'DDD' alphabetically is returned. You can then display the first five results.

The following information needs to be retained to have Previous and Next buttons that return more results:

- Previous: The title of the *first* record in the previous set, and the primary key of the *first* record in the previous set.
- Next: The title of the *first* record in the next set; the primary key of the *first* record in the next set.

These four pieces of information can be encoded into the Previous and Next buttons as in the previous part. Using your database table from the first part, write a JavaServer Page that displays the book information 20 records at a time. The page should include *Previous* and *Next* buttons to show the previous or next record set if there is one. Use the constraint query to get the Previous and Next record sets.

PROJECT-BASED EXERCISES

In this chapter, you continue the exercises from the previous chapter and create the parts of the application that reside at the middle tier and at the presentation tier. More information about these exercises and material for more exercises can be found online at

`http://www.cs.wisc.edu/~dbbook`

Exercise 7.7 Recall the Notown Records website that you worked on in Exercise 6.6. Next, you are asked to develop the actual pages for the Notown Records website. Design the part of the website that involves the presentation tier and the middle tier, and integrate the code that you wrote in Exercise 6.6 to access the database.

I. Describe in detail the set of webpages that users can access. Keep the following issues in mind:

   - All users start at a common page.
   - For each action, what input does the user provide? How will the user provide it -by clicking on a link or through an HTML form?
   - What sequence of steps does a user go through to purchase a record? Describe the high-level application flow by showing how each IIser action is handled.

2. Write the webpages in HTML without dynamic content.
3. Write a page that allows users to log on to the site. Use cookies to store the information permanently at the user's browser.
4. Augment the log-on page with JavaScript code that checks that the username consists only of the characters from a to z.
5. Augment the pages that allow users to store items in a shopping basket with a condition that checks whether the user has logged on to the site. If the user has not yet logged on, there should be no way to add items to the shopping cart. Implement this functionality using JSP by checking cookie information from the user.
6. Create the remaining pages to finish the website.

Exercise **7.8** Recall the online pharmacy project that you worked on in Exercise 6.7 in Chapter 6. Follow the analogous steps from Exercise 7.7 to design the application logic and presentation layer and finish the website.

Exercise **7.9** Recall the university database project that you worked on in Exercise 6.8 in Chapter 6. Follow the analogous steps from Exercise 7.7 to design the application logic and presentation layer and finish the website.

Exercise **7.10** Recall the airline reservation project that you worked on in Exercise 6.9 in Chapter 6. Follow the analogous steps from Exercise 7.7 to design the application logic and presentation layer and finish the website.

BIBLIOGRAPHIC NOTES

The latest version of the standards mentioned in this chapter can be found at the website of the World Wide Web Consortium (www.w3.org). It contains links to information about I-ITML, cascading style sheets, XIvIL, XSL, and much more. The book by Hall is a general introduction to Web programming technologies [357]; a good starting point on the Web is www.Webdeve1oper.com. There are many introductory books on CGI prograruning, for example [210, 198]. The JavaSoft (java.sun.com) home page is a good starting point for Servlets, .JSP, and all other Java-related technologies. The book by Hunter [394] is a good introduction to Java Servlets. Microsoft supports Active Server Pages (ASP), a comparable tedmology to .lSI'. More information about ASP can be found on the Microsoft Developer's Network horne page (msdn.microsoft.com).

There are excellent websites devoted to the advancement of XML, for example www.xm1.com and www.ibm.com/xm1. that also contain a plethora of links with information about the other standards. There are good introductory books on many diflerent aspects of XML, for example [195, 158,597,474, 381, 320]. Information about UNICODE can be found on its home page http://www.unicode.org.

Inforrnation about .lavaServer Pages ane! servlets can be found on the JavaSoft home page at java.sun.com at java.sun.com/products/jsp and at java.sun.com/products/servlet.

PART III

STORAGE AND INDEXING

8

OVERVIEW'OF STORAGE AND INDEXING

- How does a DBMS store and access persistent data?
- Why is I/O cost so important for database operations?
- How does a DBMS organize files of data records on disk to minimize I/O costs?
- What is an index, and why is it used?
- What is the relationship between a file of data records and any indexes on this file of records?
- What are important properties of indexes?
- How does a hash-based index work, and when is it most effective?
- How does a tree-based index work, and when is it most effective?
- How can we use indexes to optimize performance for a given workload?
- Key concepts: external storage, buffer manager, page I/O; file organization, heap files, sorted files; indexes, data entries, search keys, clustered index, clustered file, primary index; index organization, hash-based and tree-based indexes; cost comparison, file organizations and common operations; performance tuning, workload, composite search keys, use of clustering,

If you don't find it in the index, look very carefully through the entire catalog.

--Sears, Roebuck, and Co., Consumers' Guide, 1897

The basic abstraction of data in a DBMS is a collection of records, or a *file,* and each file consists of one or more pages. The *files and access methods*

software layer organizes data carefully to support fast access to desired subsets of records. Understanding how records are organized is essential to using a database system effectively, and it is the main topic of this chapter.

A file organization is a method of arranging the records in a file when the file is stored on disk. Each file organization makes certain operations efficient but other operations expensive.

Consider a file of employee records, each containing *age, name,* and *sal* fields, which we use as a running example in this chapter. If we want to retrieve employee records in order of increasing age, sorting the file by age is a good file organization, but the sort order is expensive to maintain if the file is frequently modified. Further, we are often interested in supporting more than one operation on a given collection of records. In our example, we may also want to retrieve all employees who make more than $5000. We have to scan the entire file to find such employee records.

A technique called *indexing* can help when we have to access a collection of records in multiple ways, in addition to efficiently supporting various kinds of selection. Section 8.2 introduces indexing, an important aspect of file organization in a DBMS. We present an overview of index data structures in Section 8.3; a more detailed discussion is included in Chapters 10 and 11.

We illustrate the importance of choosing an appropriate file organization in Section 8.4 through a simplified analysis of several alternative file organizations. The cost model used in this analysis, presented in Section 8.4.1, is used in later chapters as well. In Section 8.5, we highlight some important choices to be made in creating indexes. Choosing a good collection of indexes to build is arguably the single most powerful tool a database administrator has for improving performance.

8.1 DATA ON EXTERNAL STORAGE

A DBMS stores vast quantities of data, and the data must persist across program executions. Therefore, data is stored on external storage devices such as disks and tapes, and fetched into main memory as needed for processing. The unit of information read from or written to disk is a *page*. The size of a page is a DBMS parameter, and typical values are 4KB or 8KB.

The cost of page I/O (*input* from disk to main Inemory and *output* from memory to disk) dominates the cost of typical database operations, and database systems are carefully optimized to rninimize this cost. While the details of how

files of records are physically stored on disk and how main memory is utilized are covered in Chapter 9, the following points are important to keep in mind:

- Disks are the most important external storage devices. They allow us to retrieve any page at a (more or less) fixed cost per page. However, if we read several pages in the order that they are stored physically, the cost can be much less than the cost of reading the same pages in a random order.
- Tapes are sequential access devices and force us to read data one page after the other. They are mostly used to archive data that is not needed on a regular basis.
- Each record in a file has a unique identifier called a record id, or **rid** for short. An rid has the property that we can identify the disk address of the page containing the record by using the rid.

Data is read into memory for processing, and written to disk for persistent storage, by a layer of software called the *buffer manager*. When the *files and access methods* layer (which we often refer to as just the file layer) needs to process a page, it asks the buffer manager to fetch the page, specifying the page's rid. The buffer manager fetches the page from disk if it is not already in memory.

Space on disk is managed by the *disk space m,anager,* according to the DBMS software architecture described in Section 1.8. When the files and access methods layer needs additional space to hold new records in a file, it asks the disk space manager to allocate an additional disk page for the file; it also informs the disk space manager when it no longer needs one of its disk pages. The disk space manager keeps track of the pages in use by the file layer; if a page is freed by the file layer, the space rnanager tracks this, and reuses the space if the file layer requests a new page later on.

In the rest of this chapter, we focus on the files and access methods layer.

8.2 FILE ORGANIZATIONS AND INDEXING

The file of records is an important abstraction in a DBMS, and is implemented by the files and access methods layer of the code. A file can be created, destroyed, and have records inserted into and deleted from it. It also supports scallS; a scan operation allows us to step through all the records in the file one at a time. A relatioll is typically stored as a file of records.

The file layer stores the records in a file in a collection of disk pages. It keeps track of pages allocated to each file, and as records are inserted into and deleted from the file, it also tracks availa.ble space within pages allocated to the file.

The simplest file structure is an unordered file, or heap file. Records in a heap file are stored in random order across the pages of the file. A heap file organization supports retrieval of all records, or retrieval of a particular record specified by its rid; the file manager must keep track of the pages allocated for the file. (We defer the details of how a heap file is implemented to Chapter 9.)

An index is a data structure that organizes data records on disk to optimize certain kinds of retrieval operations. An index allows us to efficiently retrieve all records that satisfy search conditions on the search key fields of the index. We can also create additional indexes on a given collection of data records, each with a different search key, to speed up search operations that are not efficiently supported by the file organization used to store the data records.

Consider our example of employee records. We can store the records in a file organized as an index on employee age; this is an alternative to sorting the file by age. Additionally, we can create an auxiliary index file based on salary, to speed up queries involving salary. The first file contains employee records, and the second contains records that allow us to locate employee records satisfying a query on salary.

We use the term data entry to refer to the records stored in an index file. A data entry with search key value k, denoted as $k*$, contains enough information to locate (one or more) data records with search key value k. We can efficiently search an index to find the desired data entries, and then use these to obtain data records (if these are distinct from data entries).

There are three main alternatives for what to store as a data entry in an index:

1. A data entry $k*$ is an actual data record (with search key value *k*).
2. A data entry is a *(k, rid)* pair, where *rid* is the record id of a data record with search key value *k*.
3. A data entry is a $\langle k, rid\text{-}list \rangle$ pair, where *rid-list* is a list of record ids of data records with search key value *k*.

Of course, if the index is used to store actual data records, Alternative (1), each entry $k*$ is a data record with search key value *k*. We can think of such an index as a special file organization. Such an indexed file organization can be used instead of, for exarnple, a sorted file or an unordered file of records.

Alternatives (2) and (3), which contain data entries that point to data records, are independent of the file organization that is used for the indexed file (i.e.,

the file that contains the data records). Alternative (3) offers better space utilization than Alternative (2), but data entries are variable in length, depending on the number of data records with a given search key value.

If we want to build more than one index on a collection of data records-for example, we want to build indexes on both the *age* and the *sal* fields for a collection of employee records—at most one of the indexes should use Alternative (1) because we should avoid storing data records multiple times.

8.2.1 Clustered Indexes

When a file is organized so that the ordering of data records is the same as or close to the ordering of data entries in some index, we say that the index is clustered; otherwise, it clustered is an unclustered index. An index that uses Alternative (1) is clustered, by definition. An index that uses Alternative (2) or (3) can be a clustered index only if the data records are sorted on the search key field. Otherwise, the order of the data records is random, defined purely by their physical order, and there is no reasonable way to arrange the data entries in the index in the same order.

In practice, files are rarely kept sorted since this is too expensive to maintain when the data is updated. So, in practice, a clustered index is an index that uses Alternative (1), and indexes that use Alternatives (2) or (3) are unclustered. We sometimes refer to an index using Alternative (1) as a clustered file, because the data entries are actual data records, and the index is therefore a file of data records. (As observed earlier, searches and scans on an index return only its data entries, even if it contains additional information to organize the data entries.)

The cost of using an index to answer a range search query can vary tremendously based on whether the index is clustered. If the index is clustered, i.e., we are using the search key of a clustered file, the rids in qualifying data entries point to a contiguous collection of records, and we need to retrieve only a few data pages. If the index is unclustered, each qualifying data entry could contain a rid that points to a distinct data page, leading to as many data page I/Os as the number of data entries that match the range selection, as illustrated in Figure 8.1. This point is discussed further in Chapter 13.

8.2.2 Primary and Secondary Indexes

An index on a set of fields that includes the *primary key* (see Chapter 3) is called a primary index; other indexes are called secondary indexes. (The terms *primary inde.T* and *secondaTy index* are sometimes used with a different

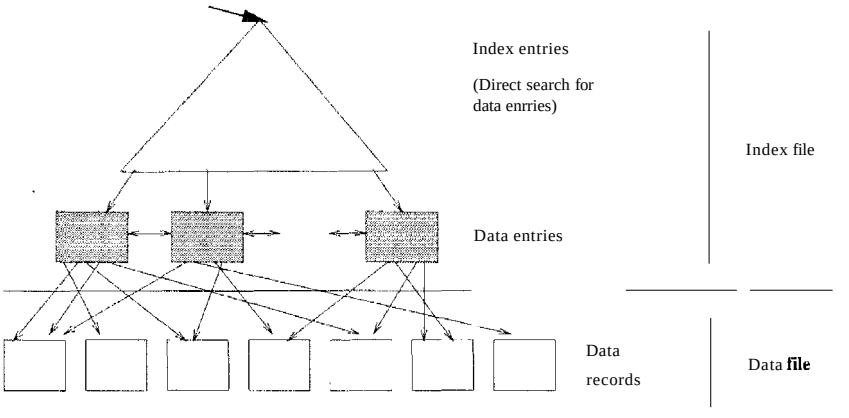


Figure 8.1 Uuelllst.ered Index Using Alternative (2)

meaning: An index that uses Alternative (1) is called a *primary index,* and one that uses Alternatives (2) or (3) is called a *secondary index.* We will be consistent with the definitions presented earlier, but the reader should be aware of this lack of standard terminology in the literature.)

Two data entries are said to be duplicates if they have the same value for the search key field associated with the index. A primary index is guaranteed not to contain duplicates, but an index on other (collections of) fields can contain duplicates. In general, a secondary index contains duplicates. If we know tha.t no duplicates exist, that is, we know that the search key contains some candidate key, we call the index a **unique** index.

An important issue is how data entries in an index are organized to support efficient retrieval of data entries.vVe discuss this next.

8.3 INDEX DATA STRUCTURES

One way to organize data entries is to hash data entries on the sea.rch key. Another way to organize data entries is to build a tree-like data structure that directs a search for data entries. We introduce these two basic approaches ill this section. We study tree-based indexing in more detail in Chapter 10 and hash-based indexing in Chapter 11.

We note that the choice of hash or tree indexing techniques can be combined with any of the three alternatives for data entries.

8.3.1 Hash-Based Indexing

We can organize records using a technique called *hashing* to quickly find records that have a given search key value. For example, if the file of employee records is hashed on the *name* field, we can retrieve all records about Joe.

In this approach, the records in a file are grouped in buckets, where a bucket consists of a primary page and, possibly, additional pages linked in a chain. The bucket to which a record belongs can be determined by applying a special function, called a hash function, to the search key. Given a bucket number, a hash-based index structure allows us to retrieve the primary page for the bucket in one or two disk l/Os.

On inserts, the record is inserted into the appropriate bucket, with 'overflow' pages allocated as necessary. To search for a record with a given search key value, we apply the hash function to identify the bucket to which such records belong and look at all pages in that bucket. If we do not have the search key value for the record, for example, the index is based on *sal* and we want records with a given age value, we have to scan all pages in the file.

In this chapter, we assume that applying the hash function to (the search key of) a record allows us to identify and retrieve the page containing the record with one I/O. In practice, hash-based index structures that adjust gracefully to inserts and deletes and allow us to retrieve the page containing a record in one to two l/Os (see Chapter 11) are known.

Hash indexing is illustrated in Figure 8.2, where the data is stored in a file that is hashed on *age;* the data entries in this first index file are the actual data records. Applying the hash function to the age field identifies the page that the record belongs to. The hash function *h* for this example is quite simple; it converts the search key value to its binary representation and uses the two least significant bits as the bucket identifier.

Figure 8.2 also shows an index with search key *sal* that contains *(sal, rid)* pairs as data entries. The *rid* (short for *record id)* component of a data entry in this second index is a pointer to a record with search key value *sal* (and is shown in the figure as an arrow pointing to the data record).

Using the terminology introduced in Section 8.2, Figure 8.2 illustrates Alternatives (1) and (2) for data entries. The file of employee records is hashed on *age,* and Alternative (1) is used for for data entries. The second index, on *sal,* also uses hashing to locate data entries, which are now *(sal, rid of employee record)* pairs; that is, Alternative (2) is used for data entries.

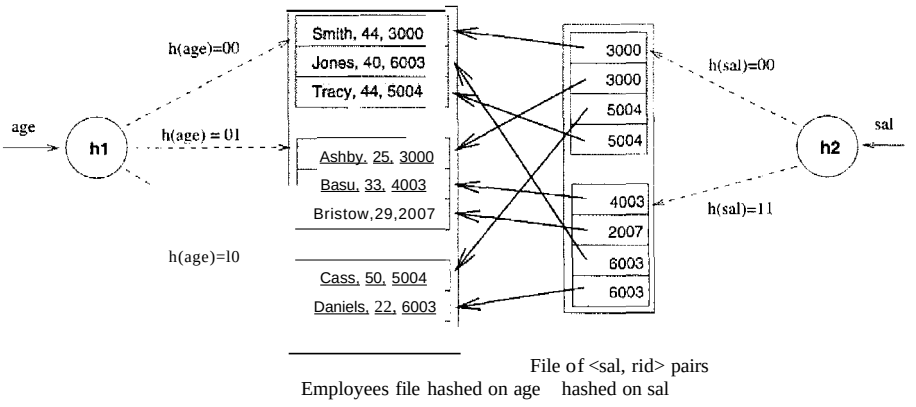


Figure 8.2 Index-Organized File Hashed on *age,* with Auxiliary Index on *sal*

Note that the search key for an index can be any sequence of one or more fields, and it need not uniquely identify records. For example, in the salary index, two data entries have the same search key value 6003. (There is an unfortunate overloading of the term *key* in the database literature. A *primary key* or *candidate* key-fields that uniquely identify a record; see Chapter 3—is unrelated to the concept of a search key.)

8.3.2 Tree-Based Indexing

An alternative to hash-based indexing is to organize records using a tree-like data structure. The data entries are arranged in sorted order by search key value, and a hierarchical search data structure is maintained that directs searches to the correct page of data entries.

Figure 8.3 shows the employee records from Figure 8.2, this time organized in a tree-structured index with search *keyage*. Each node in this figure (e.g., nodes labeled A, B, L1, L2) is a physical page, and retrieving a node involves a disk I/O.

The lowest level of the tree, called the **leaf level,** contains the data entries; in our example, these are employee records. To illustrate the ideas better, we have drawn Figure 8.3 as if there were additional employee records, some with age less than 22 and some with age greater than 50 (the lowest and highest age values that appear in Figure 8.2). Additional records with age less than 22 would appear in leaf pages to the left page L1, and records with age greater than 50 would appear in leaf pages to the right of page L3.

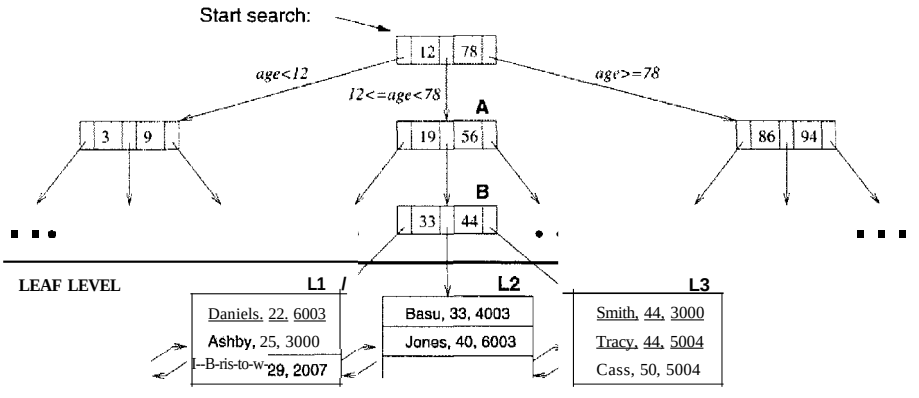


Figure 8.3 Tree·Structured Index

This structure allows us to efficiently locate all data entries with search key values in a desired range. All searches begin at the topmost node, called the root, and the contents of pages in non-leaf levels direct searches to the correct leaf page. Non-leaf pages contain node pointers separated by search key values. The node pointer to the left of a key value *k* points to a subtree that contains only data entries less than *k*. The node pointer to the right of a key value *k* points to a subtree that contains only data entries greater than or equal to *k*.

In our example, suppose we want to find all data entries with $24 < age < 50$. Each edge from the root node to a child node in Figure 8.2 has a label that explains what the corresponding subtree contains. (Although the labels for the remaining edges in the figure are not shown, they should be easy to deduce.) In our example search, we look for data entries with search key value > 24, and get directed to the middle child, node A. Again, examining the contents of this node, we are directed to node B. Examining the contents of node B, we are directed to leaf node L1, which contains data entries we are looking for.

Observe that leaf nodes L2 and L3 also contain data entries that satisfy our search criterion. To facilitate retrieval of such qualifying entries during search, all leaf pages are maintained in a doubly-linked list. Thus, we can fetch page L2 using the 'next' pointer on page L1, and then fetch page L3 using the 'next' pointer on L2.

Thus, the number of disk I/Os incurred during a search is equal to the length of a path from the root to a leaf, plus the number of leaf pages with qualifying data entries. The B+ tree is an index structure that ensures that all paths from the root to a leaf in a given tree are of the same length, that is, the structure is always balanced in height. Finding the correct leaf page is faster

than binary search of the pages in a sorted file because each non-leaf node can accommodate a very large number of node-pointers, and the height of the tree is rarely more than three or four in practice. The **height** of a balanced tree is the length of a path from root to leaf; in Figure 8.3, the height is three. The number of I/Os to retrieve a desired leaf page is four, including the root and the leaf page. (In practice, the root is typically in the buffer pool because it is frequently accessed, and we really incur just three I/Os for a tree of height three.)

The average number of children for a non-leaf node is called the **fan-out** of the tree. If every non-leaf node has n children, a tree of height h has n^h leaf pages. In practice, nodes do not have the same number of children, but using the average value F for n, we still get a good approximation to the number of leaf pages, F^h. In practice, F is at least 100, which means a tree of height four contains 100 million leaf pages. Thus, we can search a file with 100 million leaf pages and find the page we want using four I/Os; in contrast, binary search of the same file would take $\log_2 100{,}000{,}000$ (over 25) I/Os.

8.4 COMPARISON OF FILE ORGANIZATIONS

We now compare the costs of some simple operations for several basic file organizations on a collection of employee records. We assume that the files and indexes are organized according to the composite search key $\langle age, sal \rangle$, and that all selection operations are specified on these fields. The organizations that we consider are the following:

- File of randomly ordered employee records, or heap file.
- File of employee records sorted on *(age, sal)*.
- Clustered B+ tree file with search key *(age, sal)*.
- Heap file with an unclustered B+ tree index on *(age, sal)*.
- Heap file with an unclustered hash index on *(age, sal)*.

Our goal is to emphasize the importance of the choice of an appropriate file organization, and the above list includes the main alternatives to consider in practice. Obviously, we can keep the records unsorted or sort them. We can also choose to build an index on the data file. Note that even if the data file is sorted, an index whose search key differs from the sort order behaves like an index on a heap file!

The operations we consider are these:

- Scan: Fetch all records in the file. The pages in the file must be fetched from disk into the buffer pool. There is also a CPU overhead per record for locating the record on the page (in the pool).
- Search with Equality Selection: Fetch all records that satisfy an equality selection; for example, "Find the employee record for the employee with *age* 23 and *sal* 50." Pages that contain qualifying records must be fetched from disk, and qualifying records must be located within retrieved pages.
- Search with Range Selection: Fetch all records that satisfy a range selection; for example, "Find all employee records with *age* greater than 35."
- Insert a Record: Insert a given record into the file. We must identify the page in the file into which the new record must be inserted, fetch that page from disk, modify it to include the new record, and then write back the modified page. Depending on the file organization, we may have to fetch, modify, and write back other pages as well.
- Delete a Record: Delete a record that is specified using its rid. We must identify the page that contains the record, fetch it from disk, modify it, and write it back. Depending on the file organization, we may have to fetch, modify, and write back other pages as well.

8.4.1 Cost Model

In our comparison of file organizations, and in later chapters, we use a simple cost model that allows us to estimate the cost (in terms of execution time) of different database operations. We use B to denote the number of data pages when records are packed onto pages with no wasted space, and R to denote the number of records per page. The average time to read or write a disk page is D, and the average time to process a record (e.g., to compare a field value to a selection constant) is C. In the hashed file organization, we use a function, called a *hash function*, to map a record into a range of numbers; the time required to apply the hash function to a record is H. For tree indexes, we will use F to denote the fan-out, which typically is at least 100 as mentioned in Section 8.3.2.

Typical values today are D = 15 milliseconds, C and H = 100 nanoseconds; we therefore expect the cost of I/O to dominate. I/O is often (even typically) the dominant component of the cost of database operations, and so considering I/O costs gives us a good first approximation to the true costs. Further, CPU speeds are steadily rising, whereas disk speeds are not increasing at a similar pace. (On the other hand, as main memory sizes increase, a much larger fraction of the needed pages are likely to fit in memory, leading to fewer I/O requests!) We

have chosen to concentrate on the I/O component of the cost model, and we assume the simple constant *C* for in-memory per-record processing cost. Bear the follO\ving observations in mind:

- Real systems must consider other aspects of cost, such as CPU costs (and network transmission costs in a distributed database).
- Even with our decision to focus on I/O costs, an accurate model would be too complex for our purposes of conveying the essential ideas in a simple way. We therefore use a simplistic model in which we just count the number of pages read from or written to disk as a measure of I/O. We ignore the important issue of **blocked** access in our analysis-typically, disk systems allow us to read a block of contiguous pages in a single I/O request. The cost is equal to the time required to *seek* the first page in the block and transfer all pages in the block. Such blocked access can be much cheaper than issuing one I/O request per page in the block, especially if these requests do not follow consecutively, because we would have an additional seek cost for each page in the block.

We discuss the implications of the cost model whenever our simplifying assumptions are likely to affect our conclusions in an important way.

8.4.2 Heap Files

Scan: The cost is $B(D+RC)$ because we must retrieve each of B pages taking time D per page, and for each page, process R records taking time C per record.

Search with Equality Selection: Suppose that we know in advance that exactly one record matches the desired equality selection, that is, the selection is specified on a candidate key. On average, we must scan half the file, assuming that the record exists and the distribution of values in the search field is uniform. For each retrieved data page, we must check all records on the page to see if it is the desired record. The cost is $0.5B(D+RC)$. If no record satisfies the selection, however, we must scan the entire file to verify this.

If the selection is not on a candidate key field (e.g., "Find employees aged 18"), we always have to scan the entire file because records with *age* = 18 could be dispersed all over the file, and we have no idea how many such records exist.

Search with Range Selection: The entire file must be scanned because qualifying records could appear anywhere in the file, and we do not know how many qualifying records exist. The cost is $B(D+RC)$.

Insert: We assume that records are always inserted at the end of the file. We must fetch the last page in the file, add the record, and write the page back. The cost is $2D + C$.

Delete: We must find the record, remove the record from the page, and write the modified page back. We assume that no attempt is made to compact the file to reclaim the free space created by deletions, for simplicity.[1] The cost is the cost of searching plus $C + D$.

We assume that the record to be deleted is specified using the record id. Since the page id can easily be obtained from the record id, we can directly read in the page. The cost of searching is therefore D.

If the record to be deleted is specified using an equality or range condition on some fields, the cost of searching is given in our discussion of equality and range selections. The cost of deletion is also affected by the number of qualifying records, since all pages containing such records must be modified.

8.4.3 Sorted Files

Scan: The cost is $B(D + RC)$ because all pages must be examined. Note that this case is no better or worse than the case of unordered files. However, the order in which records are retrieved corresponds to the sort order, that is, all records in *age* order, and for a given age, by *sal* order.

Search with Equality Selection: We assume that the equality selection matches the sort order *(age, sal)*. In other words, we assume that a selection condition is specified on at least the first field in the composite key (e.g., *age* = 30). If not (e.g., selection *sal* = 50 or *department* = *"Toy"*), the sort order does not help us and the cost is identical to that for a heap file.

We can locate the first page containing the desired record or records, should any qualifying records exist, with a binary search in $log2B$ steps. (This analysis assumes that the pages in the sorted file are stored sequentially, and we can retrieve the ith page on the file directly in one disk I/O.) Each step requires a disk I/O and two cornparisons. Once the page is known, the first qualifying record can again be located by a binary search of the page at a cost of $Clog_2R$. The cost is $Dlog_2B + Clog_2R$, which is a significant improvement over searching heap files.

]In practice, a directory or other data structure is used to keep track of free space, and records are inserted into the first available free slot. as discussed in Chapter 9. This increases the cost of insertion and deletion a little, but not enough to affect our comparison.

If several records qualify (e.g., "Find all employees aged 18"), they are guaranteed to be adjacent to each other due to the sorting on *age*, and so the cost of retrieving all such records is the cost of locating the first such record $(D log_2 B + C log_2 R)$ plus the cost of reading all the qualifying records in sequential order. Typically, all qualifying records fit on a single page. If no records qualify, this is established by the search for the first qualifying record, which finds the page that would have contained a qualifying record, had one existed, and searches that page.

Search with Range Selection: Again assuming that the range selection matches the composite key, the first record that satisfies the selection is located as for search with equality. Subsequently, data pages are sequentially retrieved until a record is found that does not satisfy the range selection; this is similar to an equality search with many qualifying records.

The cost is the cost of search plus the cost of retrieving the set of records that satisfy the search. The cost of the search includes the cost of fetching the first page containing qualifying, or matching, records. For small range selections, all qualifying records appear on this page. For larger range selections, we have to fetch additional pages containing matching records.

Insert: To insert a record while preserving the sort order, we must first find the correct position in the file, add the record, and then fetch and rewrite all subsequent pages (because all the old records are shifted by one slot, assuming that the file has no empty slots). On average, we can assume that the inserted record belongs in the middle of the file. Therefore, we must read the latter half of the file and then write it back after adding the new record. The cost is that of searching to find the position of the new record plus $2 * (0.5B(D + RC))$, that is, search cost plus $B(D + RC)$.

Delete: We must search for the record, remove the record from the page, and write the modified page back. We must also read and write all subsequent pages because all records that follow the deleted record must be moved up to compact the free space.[2] The cost is the same as for an insert, that is, search cost plus $B(D + RC)$. Given the rid of the record to delete, we can fetch the page containing the record directly.

If records to be deleted are specified by an equality or range condition, the cost of deletion depends on the number of qualifying records. If the condition is specified on the sort field, qualifying records are guaranteed to be contiguous, and the first qualifying record can be located using binary search.

2Unlike a heap file, there is no inexpensive way to manage free space, so we account for the cost of compacting a file when a record is deleted.

8.4.4 Clustered Files

In a clustered file, extensive empirical study has shown that pages are usually at about 67 percent occupancy. Thus, the Humber of physical data pages is about *1.5B,* and we use this observation in the following analysis.

Scan: The cost of a scan is *1.5B(D + RC)* because all data pages must be examined; this is similar to sorted files, with the obvious adjustment for the increased number of data pages. Note that our cost metric does not capture potential differences in cost due to sequential I/O. We would expect sorted files to be superior in this regard, although a clustered file using ISAM (rather than B+ trees) would be close.

Search with Equality Selection: We assume that the equality selection matches the search key *(age, sal).* We can locate the first page containing the desired record or records, should any qualifying records exist, in *logF1.5B* steps, that is, by fetching all pages from the root to the appropriate leaf. In practice, the root page is likely to be in the buffer pool and we save an I/O, but we ignore this in our simplified analysis. Each step requires a disk I/O and two comparisons. Once the page is known, the first qualifying record can again be located by a binary search of the page at a cost of *Clog2R.* The cost is $DlogF1.5B + Clog_2R$, which is a significant improvement over searching even sorted files.

If several records qualify (e.g., "Find all employees aged 18"), they are guaranteed to be adjacent to each other due to the sorting on *age,* and so the cost of retrieving all such records is the cost of locating the first such record $(Dlogp1.5B + Clog_2R)$ plus the cost of reading all the qualifying records in sequential order.

Search with Range Selection: Again assuming that the range selection matches the composite key, the first record that satisfies the selection is located as it is for search with equality. Subsequently, data pages are sequentially retrieved (using the next and previous links at the leaf level) until a record is found that does not satisfy the range selection; this is similar to an equality search with many qualifying records.

Insert: To insert a record, we must first find the correct leaf page in the index, reading every page from root to leaf. Then, we must add the llew record. Most of the time, the leaf page has sufficient space for the new record, and all we need to do is to write out the modified leaf page. Occasionally, the leaf is full and we need to retrieve and modify other pages, but this is sufficiently rare

that we can ignore it in this simplified analysis. The cost is therefore the cost of search plus one write, $Dlog_F L5B + Clog_2 R + D$.

Delete: We must search for the record, remove the record from the page, and write the modified page back. The discussion and cost analysis for insert applies here as well.

8.4.5 Heap File with Unclustered Tree Index

The number of leaf pages in an index depends on the size of a data entry. We assume that each data entry in the index is a tenth the size of an employee data record, which is typical. The number of leaf pages in the index is o.1(L5B) = $O.15B$, if we take into account the 67 percent occupancy of index pages. Similarly, the number of data entries on a page $10(0.67R) = 6.7R$, taking into account the relative size and occupancy.

Scan: Consider Figure 8.1, which illustrates an unclustered index. To do a full scan of the file of employee records, we can scan the leaf level of the index and for each data entry, fetch the corresponding data record from the underlying file, obtaining data records in the sort order *(age, sal)*.

We can read all data entries at a cost of $O.15B(D + 6.7RC)$ l/Os. Now comes the expensive part: We have to fetch the employee record for each data entry in the index. The cost of fetching the employee records is one I/O per record, since the index is unclustered and each data entry on a leaf page of the index could point to a different page in the employee file. The cost of this step is $BR(D + C)$, which is prohibitively high. If we want the employee records in sorted order, we would be better off ignoring the index and scanning the employee file directly, and then sorting it. A simple rule of thumb is that a file can be sorted by a two-pass algorithm in which each pass requires reading and writing the entire file. Thus, the I/O cost of sorting a file with B pages is $4B$, which is much less than the cost of using an unclustered index.

Search with Equality Selection: We assume that the equality selection matches the sort order $\langle age, sal \rangle$. We can locate the first page containing the desired data entry or entries, should any qualifying entries exist, in $lagrO.15B$ steps, that is, by fetching all pages from the root to the appropriate leaf. Each step requires a disk I/O and two comparisons. Once the page is known, the first qua1ifying data entry can again be located by a binary search of the page at a cost of $Clog_2 6.7R$. The first qualifying data record can be fetched fronl the employee file with another I/O. The cost is $DlogpO.15B + Clag26.7R + D$, which is a significant improvement over searching sorted files.

If several records qualify (e.g., "Find all employees aged 18"), they are *not* guaranteed to be adjacent to each other. The cost of retrieving all such records is the cost oflocating the first qualifying data entry *(Dlo9pO.15B + Clo926.7R)* plus one I/O per qualifying record. The cost of using an unclustered index is therefore very dependent on the number of qualifying records.

Search with Range Selection: Again assuming that the range selection matches the composite key, the first record that satisfies the selection is located as it is for search with equality. Subsequently, data entries are sequentially retrieved (using the next and previous links at the leaf level of the index) until a data entry is found that does not satisfy the range selection. For each qualifying data entry, we incur one I/O to fetch the corresponding employee records. The cost can quickly become prohibitive as the number of records that satisfy the range selection increases. As a rule of thumb, if 10 percent of data records satisfy the selection condition, we are better off retrieving all employee records, sorting them, and then retaining those that satisfy the selection.

Insert: We must first insert the record in the employee heap file, at a cost of *2D* + C. In addition, we must insert the corresponding data entry in the index. Finding the right leaf page costs *Dl09pO.15B + Cl0926.7R*, and writing it out after adding the new data entry costs another *D.*

Delete: We need to locate the data record in the employee file and the data entry in the index, and this search step costs *Dl09FO.15B + Cl0926.7R + D*. Now, we need to write out the modified pages in the index and the data file, at a cost of *2D*.

8.4.6 Heap File With Unclustered Hash Index

As for unclustered tree indexes, we assume that each data entry is one tenth the size of a data record. We consider only static hashing in our analysis, and for simplicity we assume that there are no overflow chains.[3]

In a static hashed file, pages are kept at about SO percent occupancy (to leave space for future insertions and minimize overflows as the file expands). This is achieved by adding a new page to a bucket when each existing page is SO percent full, when records are initially loaded into a hashed file structure. The number of pages required to store data entries is therefore 1.2.5 times the number of pages when the entries are densely packed, that is, *1.25(0.10B) = O.125B*. The number of data entries that fit on a page is 1O(O.80R) = *8R*, taking into account the relative size and occupancy.

[3]The dynamic variants of hashing are less susceptible to the problem of overflow chains, and have a slight.ly higher average cost per search, but are otherwise similar to the static version.

Scan: As for an unclustered tree index, all data entries can be retrieved inexpensively, at a cost of O.125B(D + *8RC)* I/Os. However, for each entry, we incur the additional cost of one I/O to fetch the corresponding data record; the cost of this step is *BR(D + C)*. This is prohibitively expensive, and further, results are unordered. So no one ever scans a hash index.

Search with Equality Selection: This operation is supported very efficiently for matching selections, that is, equality conditions are specified for each field in the composite search key *(age, sal)*. The cost of identifying the page that contains qualifying data entries is *H*. Assuming that this bucket consists of just one page (i.e., no overflow pages), retrieving it costs *D*. If we assume that we find the data entry after scanning half the records on the page, the cost of scanning the page is *O.5(8R)C = 4RC*. Finally, we have to fetch the data record from the employee file, which is another *D*. The total cost is therefore *H + 2D + 4RC*, which is even lower than the cost for a tree index.

If several records qualify, they are *not* guaranteed to be adjacent to each other. The cost of retrieving all such records is the cost of locating the first qualifying data entry *(H + D + 4RC)* plus one I/O per qualifying record. The cost of using an unclustered index therefore depends heavily on the number of qualifying records.

Search with Range Selection: The hash structure offers no help, and the entire heap file of employee records must be scanned at a cost of *B(D + RC)*.

Insert: We must first insert the record in the employee heap file, at a cost of *2D* + C. In addition, the appropriate page in the index must be located, modified to insert a new data entry, and then written back. The additional cost is *H + 2D* + C.

Delete: We need to locate the data record in the employee file and the data entry in the index; this search step costs *H + 2D + 4RC*. Now, we need to write out the modified pages in the index and the data file, at a cost of *2D*.

8.4.7 Comparison of I/O Costs

Figure 8.4 compares I/O costs for the various file organizations that we discussed. A heap file has good storage efficiency and supports fast scanning and insertion of records. However, it is slow for searches and deletions.

A sorted file also offers good storage efficiency. but insertion and deletion of records is slow. Searches are fa.ster than in heap files. It is worth noting that, in a real DBMS, a file is almost never kept fully sorted.

File Type	*Scan*	*Equality Search*	*Range Search*	*Insert*	*Delete*
Heap	BD	$0.5BD$	BD	$2D$	*Search+* D
Sorted	BD	$Dlog_2B$	*Dlog2B +# matching pages*	*Search* + BD	*Search+ BD*
Clustered	$1.5BD$	*DlogF1.5B*	*Dlo9F1.5B+# matching pages*	*Search* + D	*Search+* D
Unclustered tree index	$BD(R$ + 0.15)	D(1 + *logFO.15B)*	*D(lo9FO.15B+# matching records)*	*D(3* + *logFO.15B)*	*Search+* 2D
Unclustered hash index	$BD(R$ + 0.125)	2D	*BD*	4D	*Search+* 2D

Figure 8.4 A Comparison of I/O Costs

A clustered file offers all the advantages of a sorted file *and* supports inserts and deletes efficiently. (There is a space overhead for these benefits, relative to a sorted file, but the trade-off is well worth it.) Searches are even faster than in sorted files, although a sorted file can be faster when a large number of records are retrieved sequentially, because of blocked I/O efficiencies.

Unclustered tree and hash indexes offer fast searches, insertion, and deletion, but scans and range searches with many matches are slow. Hash indexes are a little faster on equality searches, but they do not support range searches.

In summary, Figure 8.4 demonstrates that no one file organization is uniformly superior in all situations.

8.5 INDEXES AND PERFORMANCE TUNING

In this section, we present an overview of choices that arise when using indexes to improve performance in a database system. The choice of indexes has a tremendous impact on system performance, and must be made in the context of the expected workload, or typical mix of queries and update operations.

A full discussion of indexes and performance requires an understanding of database query evaluation and concurrency control. We therefore return to this topic in Chapter 20, where we build on the discussion in this section. In particular, we discuss examples involving multiple tables in Chapter 20 because they require an understanding of join algorithms and query evaluation plans.

8.5.1 Impact of the Workload

The first thing to consider is the expected workload and the common operations. Different file organizations and indexes, as we have seen, support different operations well.

In generaL an index supports efficient retrieval of data entries that satisfy a given selection condition. Recall from the previous section that there are two important kinds of selections: equality selection and range selection. Hash-based indexing techniques are optimized only for equality selections and fare poorly on range selections. where they are typically worse than scanning the entire file of records. Tree-based indexing techniques support both kinds of selection conditions efficiently, explaining their widespread use.

Both tree and hash indexes can support inserts, deletes, and updates quite efficiently. Tree-based indexes, in particular, offer a superior alternative to maintaining fully sorted files of records. In contrast to simply maintaining the data entries in a sorted file, our discussion of (B+ tree) tree-structured indexes in Section 8.3.2 highlights two important advantages over sorted files:

1. We can handle inserts and deletes of data entries efficiently.
2. Finding the correct leaf page when searching for a record by search key value is much faster than binary search of the pages in a sorted file.

The one relative disadvantage is that the pages in a sorted file can be allocated in physical order on disk, making it much faster to retrieve several pages in sequential order. Of course. inserts and deletes on a sorted file are extremely expensive. A variant of B+ trees, called Indexed Sequential Access Method (ISAM), offers the benefit of sequential allocation of leaf pages, plus the benefit of fast searches. Inserts and deletes are not handled as well as in B+ trees, but are rnuch better than in a sorted file. \Ve will study tree-structured indexing in detail in Chapter 10.

8.5.2 Clustered Index Organization

As we saw in Section 8.2.1, a clustered index is really a file organization for the underlying data records. Data records can be la.rge, and we should avoid replicating them; so there can be at most one clustered index on a given collection of records. On the other hand, we can build several uncIustered indexes on a data file. Suppose that employee records are sorted by *age*, or stored in a clustered file with search keyage. If. in addition. we have an index on the *sal* field, the latter nlUst be an llnclllstered index. We can also build an uncIustered index on. say, *department*, if there is such a field.

Clustered indexes, while less expensive to maintain than a fully sorted file, are nonetheless expensive to maintain. When a new record has to be inserted into a full leaf page, a new leaf page must be allocated and sorne existing records have to be moved to the new page. If records are identified by a combination of page id and slot, as is typically the case in current database systems, all places in the database that point to a moved record (typically, entries in other indexes for the same collection of records) must also be updated to point to the new location. Locating all such places and making these additional updates can involve several disk I/Os. Clustering must be used sparingly and only when justified by frequent queries that benefit from clustering. In particular, there is no good reason to build a clustered file using hashing, since range queries cannot be answered using hash-indexes.

In dealing with the limitation that at most one index can be clustered, it is often useful to consider whether the information in an index's search key is sufficient to answer the query. If so, modern database systems are intelligent enough to avoid fetching the actual data records. For example, if we have an index on *age,* and we want to compute the average age of employees, the DBMS can do this by simply examining the data entries in the index. This is an example of an index-only evaluation. In an index-only evaluation of a query we need not access the data records in the files that contain the relations in the query; we can evaluate the query completely through indexes on the files. An important benefit of index-only evaluation is that it works equally efficiently with only unclustered indexes, as only the data entries of the index are used in the queries. Thus, unclustered indexes can be used to speed up certain queries if we recognize that the DBMS will exploit index-only evaluation.

Design Examples Illustrating Clustered Indexes

To illustrate the use of a clustered index 011 a range query, consider the following example:

```
SELECT    E.dno
FROM      Employees E
WHERE     E.age > 40
```

If we have a H+ tree index on *age,* we can use it to retrieve only tuples that satisfy the selection E.*age*> 40. Whether such an index is worthwhile depends first of all on the selectivity of the condition. What fraction of the employees are older than 40'1 If virtually everyone is older than 40, we gain little by using an index 011 *age*; a sequential scan of the relation would do almost as well. However, suppose that only 10 percent of the employees are older than 40. Now, is an index useful? The answer depends on whether the index is clustered. If the

index is unclustered, we could have one page I/O per qualifying employee, and this could be more expensive than a sequential scan, even if only 10 percent of the employees qualify! On the other hand, a clustered B+ tree index on *age* requires only 10 percent of the I/Os for a sequential scan (ignoring the few I/Os needed to traverse from the root to the first retrieved leaf page and the I/Os for the relevant index leaf pages).

As another example, consider the following refinement of the previous query:

```
SELECT    Kdno, COUNT(*)
FROM      Employees E
WHERE     E.age> 10
GROUP BY  E.dno
```

If a B+ tree index is available on *age,* we could retrieve tuples using it, sort the retrieved tuples on *dna,* and so answer the query. However, this may not be a good plan if virtually all employees are more than 10 years old. This plan is especially bad if the index is not clustered.

Let us consider whether an index on *dna* might suit our purposes better. We could use the index to retrieve all tuples, grouped by *dna,* and for each *dna* count the number of tuples with $age > 10$. (This strategy can be used with both hash and B+ tree indexes; we only require the tuples to be *grouped,* not necessarily *sorted,* by *dna.)* Again, the efficiency depends crucially on whether the index is clustered. If it is, this plan is likely to be the best if the condition on *age* is not very selective. (Even if we have a clustered index on *age,* if the condition on *age* is not selective, the cost of sorting qualifying tuples on *dna* is likely to be high.) If the index is not clustered, we could perform one page I/O per tuple in Employees, and this plan would be terrible. Indeed, if the index is not clustered, the optimizer will choose the straightforward plan based on sorting on *dna.* Therefore, this query suggests that we build a clustered index on *dna* if the condition on *age* is not very selective. If the condition is very selective, we should consider building an index (not necessarily clustered) on *age* instead.

Clustering is also important for an index on a search key that does not include a candidate key, that is, an index in which several data entries can have the same key value. To illustrate this point, we present the following query:

```
SELECT  E.dno
FROM    Employees E
WHERE   E.hobby='Stamps'
```

If many people collect stamps, retrieving tuples through an unclustered index on *hobby* can be very inefficient. It may be cheaper to simply scan the relation to retrieve all tuples and to apply the selection on-the-fly to the retrieved tuples. Therefore, if such a query is important, we should consider making the index on *hobby* a clustered index. On the other hand, if we assume that *eid* is a key for Employees, and replace the condition *E.hobby='Stamps'* by *E.eid=552,* we know that at most one Employees tuple will satisfy this selection condition. In this case, there is no advantage to making the index clustered.

The next query shows how aggregate operations can influence the choice of indexes:

```
SELECT    E.dno, COUNT(*)
FROM      Employees E
GROUP BY  E.dno
```

A straightforward plan for this query is to sort Employees on *dno* to compute the count of employees for each *dno.* However, if an index-hash or B+ tree---on *dno* is available, we can answer this query by scanning only the index. For each *dno* value, we simply count the number of data entries in the index with this value for the search key. Note that it does not matter whether the index is clustered because we never retrieve tuples of Employees.

8.5.3 Composite Search Keys

The search key for an index can contain several fields; such keys are called composite search keys or concatenated keys. As an example, consider a collection of employee records, with fields *name, age,* and *sal,* stored in sorted order by *name.* Figure 8.5 illustrates the difference between a composite index with key *(age, sal⟩*, a composite index with key *(sal, age),* an index with key *age,* and an index with key *sal.* All indexes shown in the figure use Alternative (2) for data entries.

If the search key is composite, an equality query is one in which *each* field in the search key is bound to a constant. For example, we can ask to retrieve all data entries with *age* = 20 and *sal* = 10. The hashed file organization supports only equality queries, since a hash function identifies the bucket containing desired records only if a value is specified for each field in the search key.

With respect to a composite key index, in a range query not all fields in the search key are bound to constants. For example, we can ask to retrieve all data entries with *age* = 20; this query implies that any value is acceptable for the *sal* field. As another example of a range query, we can ask to retrieve all data entries with *age* < 30 and *sal> 40.*

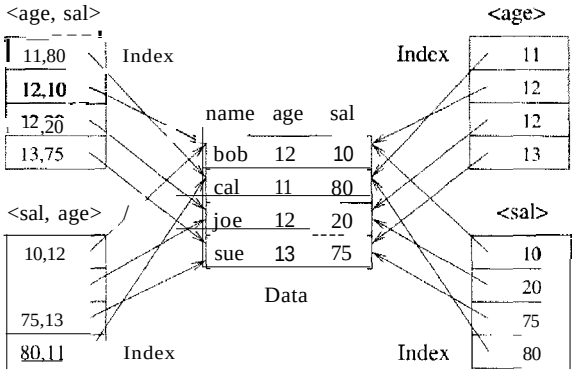


Figure 8.5 Composite Key Indexes

Nate that the index cannot help on the query *sal* > 40, because, intuitively, the index organizes records by *age* first and then *sal.* If *age* is left unspecified, qualifying records could be spread across the entire index. We say that an index **matches** a selection condition if the index can be used to retrieve just the tuples that satisf:y the condition. For selections of the form *condition* $\wedge \ldots \wedge$ *condition,* we can define when an index matches the selection as 1'01lows:[4] For a hash index, a selection matches the index if it includes an equality condition ('field = constant') on every field in the composite search key for the index. For a tree index, a selection matches the index if it includes an equality or range condition on a *prefix* of the composite search key. (As examples, *(age)* and $\langle age, sal, department)$ are prefixes of key *(age, sal,* $department\rangle$, but $\langle age, department)$ and *(sal,* $department\rangle$ are not.)

Trade-offs in Choosing Composite Keys

A composite key index can support a broader range of queries because it matches more selection conditions. Further, since data entries in a composite index contain more information about the data record (i.e., more fields than a single-attribute index), the opportunities for index-only evaluation strategies are increased. (Recall from Section 8.5.2 that an index-only evaluation does not need to access data records, but finds all required field values in the data entries of indexes.)

On the negative side, a composite index must be updated in response to any operation (insert, delete, or update) that modifies *any* field in the search key. A composite index is also likely to be larger than a single-attribute search key

[4] For a more general discussion, see Section 14.2.)

index because the size of entries is larger. For a composite B+ tree index, this also means a potential increase in the number of levels, although key COlnpression can be used to alleviate this problem (see Section 10.8.1).

Design Examples of Composite Keys

Consider the following query, which returns all employees with $20 < age < 30$ and $3000 < sal < 5000$:

```
SELECT  E.eid
FROM    Employees E
WHERE   E.age BETWEEN 20 AND 30
        AND E.sal BETWEEN 3000 AND 5000
```

A composite index on *(age, sal)* could help if the conditions in the WHERE clause are fairly selective. Obviously, a hash index will not help; a B+ tree (or ISAM) index is required. It is also clear that a clustered index is likely to be superior to an unclustered index. For this query, in which the conditions on *age* and *sal* are equally selective, a composite, clustered B+ tree index on *(age, sal)* is as effective as a composite, clustered B+ tree index on *(sal, age).* However, the order of search key attributes can sometimes make a big difference, as the next query illustrates:

```
SELECT  E.eid
FROM    Employees E
WHERE   E.age = 25
        AND E.sal BETWEEN 3000 AND 5000
```

In this query a composite, clustered B+ tree index on *(age, sal)* will give good performance because records are sorted by *age* first and then (if two records have the same *age* value) by *sal.* Thus, all records with *age* = 25 are clustered together. On the other hand, a composite, clustered B+ tree index on *(sal, age)* will not perform as well. In this case, records are sorted by *sal* first, and therefore two records with the same *age* value (in particular, with *age* = 25) may be quite far apart. In effect, this index allows us to use the range selection on *sal,* but not the equality selection on *age,* to retrieve tuples. (Good performance on both variants of the query can be achieved using a single *spatial* index. \:Ye discuss spatial indexes in Chapter 28.)

Composite indexes are also useful in dealing with many aggregate queries. Consider:

```
SELECT  AVG  (E.sal)
```

```
FROM    Employees E
WHERE   E.age = 25
        AND Ksal BETWEEN 3000 AND 5000
```

A composite B+ tree index on *(age, sal)* allows us to answer the query with an index-only scan. A composite B+ tree index on *(sal, age)* also allows us to answer the query with an index-only scan, although more index entries are retrieved in this case than with an index on *(age, sal)*.

Here is a variation of an earlier example:

```
SELECT    Kdno, COUNT(*)
FROM      Employees E
WHERE     E.sal=lO,OOO
GROUP BY  Kdno
```

An index on *dna* alone does not allow us to evaluate this query with an index-only scan, because we need to look at the *sal* field of each tuple to verify that *sal* = 10, 000. However, we can use an index-only plan if we have a composite B+ tree index on *(sal, dna)* or *(dna, sal)*. In an index with key *(sal, dno)*, all data entries with *sal* = 10,000 are arranged contiguously (whether or not the index is clustered). Further, these entries are sorted by *dna*, making it easy to obtain a count for each *dna* group. Note that we need to retrieve only data entries with *sal* = 10, 000.

It is worth observing that this strategy does not work if the WHERE clause is modified to use *sal* > 10, 000. Although it suffices to retrieve only index data entries-that is, an index-only strategy still applies-these entries must now be sorted by *dna* to identify the groups (because, for example, two entries with the same *dna* but different *sal* values may not be contiguous). An index with key *(dna, sal)* is better for this query: Data entries with a given *dna* value are stored together, and each such group of entries is itself sorted by *sal*. For each *dna* group, we can eliminate the entries with *sal* not greater than 10,000 and count the rest. (Using this index is less efficient than an index-only scan with key *(sal, dna)* for the query with *sal* = 10, 000, because we must read all data entries. Thus, the choice between these indexes is influenced by which query is more common.)

As another example, suppose we want to find the minimum *sal* for each *dna:*

```
SELECT    Kdno, MIN(E.sal)
FROM      Employees E
GROUP BY  E.dno
```

An index on *dna* alone does not allow us to evaluate this query with an index-only scan. However, we can use an index-only plan if we have a composite B+ tree index on $\langle dno, sal \rangle$. Note that all data entries in the index with a given *dna* value are stored together (whether or not the index is clustered). Further, this group of entries is itself sorted by *8al.* An index on *(sal, dna)* enables us to avoid retrieving data records, but the index data entries must be sorted on *dno.*

8.5.4 Index Specification in SQL: 1999

A natural question to ask at this point is how we can create indexes using SQL. The SQL:1999 standard does *not* include any statement for creating or dropping index structures. In fact, th.e standard does not even require SQL implementations to support indexes! In practice, of course, every commercial relational DBMS supports one or more kinds of indexes. The following command to create a B+ tree index-we discuss B+ tree indexes in Chapter 10—is illustrative:

```
CREATE INDEX IndAgeRating ON Students
      WITH STRUCTURE = BTREE,
           KEY = (age, gpa)
```

This specifies that a B+ tree index is to be created on the Students table using the concatenation of the *age* and *gpa* columns as the key. Thus, key values are pairs of the form *(age, gpa)*, and there is a distinct entry for each such pair. Once created, the index is automatically maintained by the DBMS adding or removing data entries in response to inserts or deletes of records on the Students relation.

8.6 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- 'Where does a DBMS store persistent data? How does it bring data into main memory for processing? What DBMS component reads and writes data from main memory, and what is the unit of I/O? (Section 8.1)
- 'What is a *file organization?* What is an *index?* What is the relationship between files and indexes? Can we have several indexes on a single file of records? Can an index itself store data records (i.e., act as a file)? (Section 8.2)
- What is the *search key* for an index? What is a *data entry* in an index? (Section 8.2)

- What is a *clustered* index? What is a *primary index?* How many clustered indexes can you build on a file? How many unclustered indexes can you build? **(Section 8.2.1)**
- How is data organized in a hash-based index? \Vhen would you use a hash-based index? **(Section 8.3.1)**
- How is data organized in a tree-based index? When would you use a tree-based index? **(Section 8.3.2)**
- Consider the following operations: *scans, equality and range selections, inserts, and deletes,* and the following file organizations: *heap files, sorted files, clustered files, heap files with an unclustered tree index on the search key,* and *heap files with an unclusteTed hash index.* Which file organization is best suited for each operation? **(Section 8.4)**
- What are the main contributors to the cost of database operations? Discuss a simple cost model that reflects this. **(Section 8.4.1)**
- How does the expected workload influence physical database design decisiems such as what indexes to build? Why is the choice of indexes a central aspect of physical database design? **(Section 8.5)**
- What issues are considered in using clustered indexes? What is an *indcl;-only* evaluation method? What is its primary advantage? **(Section 8.5.2)**
- What is a *composite 8earch key?* What are the pros and cons of composite search keys? **(Section 8.5.3)**
- What SQL commands support index creation? **(Section 8.5.4)**

EXERCISES

Exercise 8.1 Answer the following questions about data on external storage in a DBMS:

1. \Vhy does a DBMS store data on external storage?
2. Why are I/O costs important in a DBMS?
3. What is a record id? Given a record's id, how many I/Os are needed to fetch it into main memory?
4. \Vhat is the role of the buffer manager in a DBMS? What is the role of the disk space manager? How do these layers interact with the file and access methods layer?

Exercise 8.2 Answer the following questions about files and indexes:

1. What operations arc supported by the file of records abstraction?
2. \Vhat is an index on a file of records? \Nhat is a search key for an index? Why do we need indexes?

	name		age	gpa
53831	Madayan	madayan@music	11	1.8
53832	Gulclu	guldu@music	12	2.0
53666	Jones	jones@cs	18	3.4
53688	Smith	smith@ee	19	3.2
53650	Smith	smith@math	19	3.8

Figure 8.6 An Instance of the Students Relation, Sorted by *age*

3. What alternatives are available for the data entries in an index?
4. What is the difference between a primary index and a secondary index? \Vhat is a duplicate data entry in an index? Can a primary index contain duplicates?
5. What is the difference between a clustered index and an unclustered index? If an index contains data records as 'data entries,' can it be unclustered?
6. How many clustered indexes can you create on a file? Woule! you always create at least one clustered index for a file?
7. Consider Alternatives (1), (2) and (3) for 'data entries' in an index, as discussed in Section 8.2 . Are all of them suitable for secondary indexes? Explain.

Exercise 8.3 Consider a relation stored as a randomly ordered file for which the only index is an unclustered index on a field called *sal*. If you want to retrieve all records with $sal > 20$, is using the index always the best alternative? Explain.

Exercise 8.4 Consider the instance of the Students relation shown in Figure 8.6, sorted by *age*: For the purposes of this question, assume that these tuples are stored in a sorted file in the order shown; the first tuple is on page 1 the second tuple is also on page 1; and so on. Each page can store up to three data records; so the fourth tuple is on page 2.

Explain what the data entries in each of the following indexes contain. If the order of entries is significant, say so and explain why. If such all index cannot be constructeel, say so and explain why.

1. An unclustereel index on *age* using Alternative (1).
2. An unclusterecl index on *age* using Alternative (2).
3. An unclustered index on *age* using Alternative (3).
4. A clustered index on *age* using Alternative (1).
5. A clustered index on *age* using Alt.ernative (2).
6. A clustered index on *age* using Alternative (3).
7. An unc:lustered index on *gpa* using Alternative (1).
8. An unclustered index on *gpa* using Alternative (2).
9. An unclustered index on *gpa* using Alternative (3).
10. A clustered index on *gpa* using Alternative (1).
11. A clustered index on *gpa* using Alternative (2).
12. A clustered index on *gpa* using Alternative (3).

File Type	Scan	Equality Search	Range Search	Insert	Delete
Heap file					
Sorted file					
Clustered file					
Unclustered tree index					
Unclustered hash index					

Figure 8.7 I/O Cost Comparison

Exercise 8.5 Explain the difference between Hash indexes and B+-tree indexes. In particular, discuss how equality and range searches work, using an example.

Exercise 8.6 Fill in the I/O costs in Figure 8.7.

Exercise 8.7 If you were about to create an index on a relation, what considerations would guide your choice? Discuss:

1. The choice of primary index.
2. Clustered versus unclustered indexes.
3. Hash versus tree indexes.
4. The use of a sorted file rather than a tree-based index.
5. Choice of search key for the index. What is a composite search key, and what considerations are made in choosing composite search keys? What are index-only plans, and what is the influence of potential index-only evaluation plans on the choice of search key for an index?

Exercise 8.8 Consider a delete specified using an equality condition. For each of the five file organizations, what is the cost if no record qualifies? What is the cost if the condition is not on a key?

Exercise 8.9 What main conclusions can you draw from the discussion of the five basic file organizations discussed in Section 8.4? Which of the five organizations would you choose for a file where the most frequent operations are as follows?

1. Search for records based on a range of field values.
2. Perform inserts and scans, where the order of records docs not matter.
3. Search for a record based on a particular field value.

Exercise 8.10 Consider the following relation:

Emp(*eid:* integer, *sal:* integer, *age:* real, *did:* integer)

There is a clustered index on *cid* and an llnclustered index on *age*.

1. How would you use the indexes to enforce the constraint that *eid* is a key?
2. Give an example of an update that is *definitely speeded up* because of the available indexes. (English description is sufficient.)

3. Give an example of an update that is *definitely slowed down* because of the indexes. (English description is sufficient.)
4. Can you give an example of an update that is neither speeded up nor slowed down by the indexes?

Exercise 8.11 Consider the following relations:

Emp(*eid:* integer, *ename:* varchar, *sal:* integer, *age:* integer, *did:* integer)
Dept(did: integer, *budget:* integer, *floor:* integer, *mgr_eid:* integer)

Salaries range from $10,000 to $100,000, ages vary from 20 to 80, each department has about five employees on average, there are 10 floors, and budgets vary from $10,000 to $1 million. You can assume uniform distributions of values.

For each of the following queries, which of the listed index choices would you choose to speed up the query? If your database system does not consider index-only plans (i.e., data records are always retrieved even if enough information is available in the index entry), how would your answer change? Explain briefly.

1. Query: *Print ename, age, and sal for all employees.*
   (a) Clustered hash index on *(ename, age, sal)* fields of Emp.
   (b) Unclustered hash index on *(ename, age, sal)* fields of Emp.
   (c) Clustered B+ tree index on *(ename, age, sal)* fields of Emp.
   (d) Unclustered hash index on *(eid, did)* fields of Emp.
   (e) No index.
2. Query: *Find the dids of departments that are on the 10th floor and have a budget of less than $15,000.*
   (a) Clustered hash index on the *floor* field of Dept.
   (b) Unclustered hash index on the *floor'* field of Dept.
   (c) Clustered B+ tree index on *(floor, budget)* fields of Dept.
   (d) Clustered B+ tree index on the *budget* field of Dept.
   (e) No index.

PROJECT-BASED EXERCISES

Exercise 8.12 Answer the following questions:

1. What indexing techniques are supported in Minibase?
2. What alternatives for data entries are supported'?
3. Are clustered indexes supported?

BIBLIOGRAPHIC NOTES

Several books discuss file organization in detail [29, 312, 442, 531, 648, 695, 775].

Bibliographic: notes for hash-indexes and B+-trees are included in Chapters 10 and 11.

9

STORING DATA: DISKS AND FILES

- What are the different kinds of memory in a computer system?
- What are the physical characteristics of disks and tapes, and how do they affect the design of database systems?
- What are RAID storage systems, and what are their advantages?
- How does a DBMS keep track of space on disks? How does a DBMS access and modify data on disks? What is the significance of *pages* as a unit of storage and transfer?
- How does a DBMS create and maintain files of records? How are records arranged on pages, and how are pages organized within a file?
- Key concepts: memory hierarchy, persistent storage, random versus sequential devices; physical disk architecture, disk characteristics, seek time, rotational delay, transfer time; RAID, striping, mirroring, RAID levels; disk space manager; buffer manager, buffer pool, replacement policy, prefetching, forcing; file implementation, page organization, record organization

A memory is what is left when something happens and does not cornpletely unhappen.

. Edward DeBono

This chapter initiates a study of the internals of an RDBivIS. In terms of the DBMS architecture presented in Section 1.8, it covers the disk space manager,

the buffer manager, and implementation-oriented aspects of the *Jiles and access methods* layer.

Section 9.1 introduces disks and tapes. Section 9.2 describes RAID disk systems. Section 9.3 discusses how a DBMS manages disk space, and Section 9.4 explains how a DBMS fetches data from disk into main memory. Section 9.5 discusses how a collection of pages is organized into a file and how auxiliary data structures can be built to speed up retrieval of records from a file. Section 9.6 covers different ways to arrange a collection of records on a page, and Section 9.7 covers alternative formats for storing individual records.

9.1 THE MEMORY HIERARCHY

Memory in a computer system is arranged in a hierarchy, as shown in Figure 9.1. At the top, we have primary storage, which consists of cache and main memory and provides very fast access to data. Then comes secondary storage, which consists of slower devices, such as magnetic disks. Tertiary storage is the slowest class of storage devices; for example, optical disks and tapes. Currently, the cost of a given amount of main memory is about 100 times

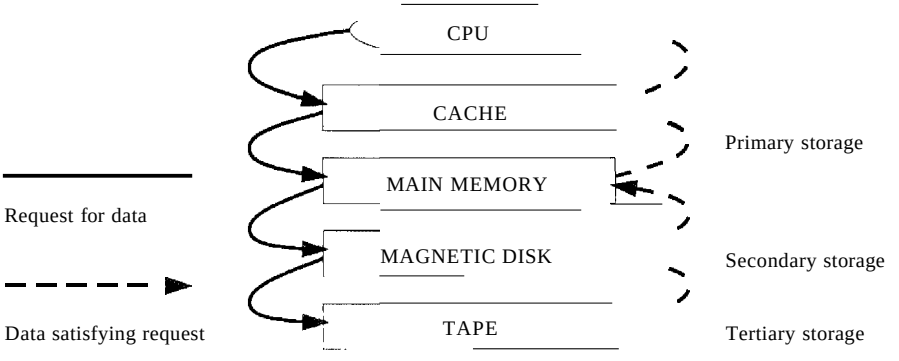


Figure 9.1 The Ivlemory Hierarchy

the cost of the same amount of disk space, and tapes are even less expensive than disks. Slower storage devices such as tapes and disks play an important role in database systems because the amount of data is typically very large. Since buying enough main memory to store all data is prohibitively expensive, we must store data on tapes and disks and build database systems that can retrieve data from lower levels of the memory hierarchy into main mernory as needed for processing.

There are reasons other than cost for storing data on secondary and tertiary storage. On systems with 32-bit addressing, only 2^{32} bytes can be directly referenced in main memory; the number of data objects may exceed this number! Further, data must be maintained across program executions. This requires storage devices that retain information when the computer is restarted (after a shutdown or a crash); we call such storage nonvolatile. Primary storage is usually volatile (although it is possible to make it nonvolatile by adding a battery backup feature), whereas secondary and tertiary storage are nonvolatile.

Tapes are relatively inexpensive and can store very large amounts of data. They are a good choice for *archival* storage, that is, when we need to maintain data for a long period but do not expect to access it very often. A Quantum DLT 4000 drive is a typical tape device; it stores 20 GB of data and can store about twice as much by compressing the data. It records data on 128 *tape tracks*, which can be thought of as a linear sequence of adjacent bytes, and supports a sustained transfer rate of 1.5 MB/sec with uncompressed data (typically 3.0 MB/sec with compressed data). A single DLT 4000 tape drive can be used to access up to seven tapes in a stacked configuration, for a maximum compressed data capacity of about 280 GB.

The main drawback of tapes is that they are sequential access devices. We must essentially step through all the data in order and cannot directly access a given location on tape. For example, to access the last byte on a tape, we would have to wind through the entire tape first. This makes tapes unsuitable for storing *operational data*, or data that is frequently accessed. Tapes are mostly used to back up operational data periodically.

9.1.1 Magnetic Disks

Magnetic disks support direct access to a desired location and are widely used for database applications. A DBMS provides seamless access to data on disk; applications need not worry about whether data is in main memory or disk. To understand how disks work, eonsider Figure 9.2, which shows the structure of a disk in simplified form.

Data is stored on disk in units called disk blocks. A disk block is a contiguous sequence of bytes and is the unit in which data is written to a disk and read from a disk. Bloc:ks are arranged in concentric rings called tracks, on one or more platters. Tracks can be recorded on one or both surfaces of a platter; we refer to platters as single-sided or double-sided, accordingly. The set of all tracks with the same diameter is called a cylinder, because the space occupied by these tracks is shaped like a cylinder; a cylinder contains one track per platter surface. Each track is divided into arcs, called sectors, whose size is a

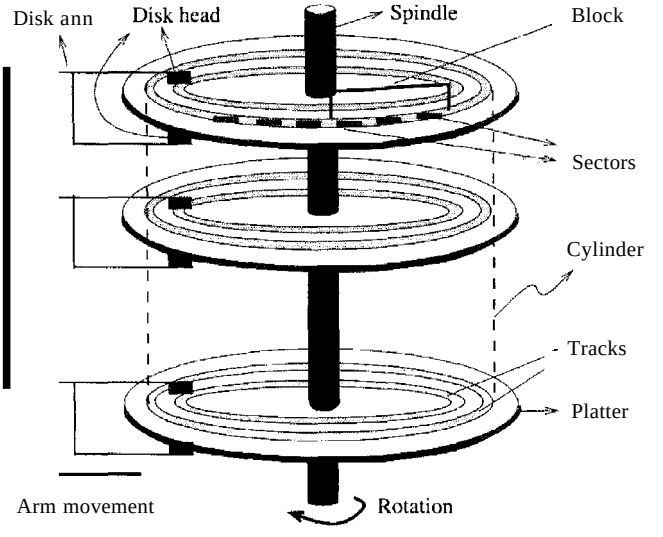


Figure 9.2 Structure of a Disk

characteristic of the disk and cannot be changed. The size of a disk block can be set when the disk is initialized as a multiple of the sector size.

An array of disk heads, one per recorded surface, is moved as a unit; when one head is positioned over a block, the other heads are in identical positions with respect to their platters. To read or write a block, a disk head must be positioned on top of the block.

Current systems typically allow at most one disk head to read or write at any one time. All the disk heads cannot read or write in parallel—this technique would increase data transfer rates by a factor equal to the number of disk heads and considerably speed up sequential scans. The reason they cannot is that it is very difficult to ensure that all the heads are perfectly aligned on the corresponding tracks. Current approaches are both expensive and more prone to faults than disks with a single active heacl. In practice, very few commercial products support this capability and then only in a limited way; for example, two disk heads may be able to operate in parallel.

A disk controller interfaces a disk drive to the computer. It implements commands to read or write a sector by moving the arm assembly and transferring data to and from the disk surfaces. A checksum is computed for when data is written to a sector and stored with the sector. The checksum is computed again when the data on the sector is read back. If the sector is corrupted or the

An Example of a Current Disk: The IBM Deskstar 14GPX. The IBM Deskstar 14GPX is a 3.5 inch, 14.4 GB hard disk with an average seek time of 9.1 milliseconds (msec) and an average rotational delay of 4.17 msec. However, the time to seek from one track to the next is just 2.2 msec, the maximum seek time is 15.5 msec. The disk has five double-sided platters that spin at 7200 rotations per minute. Each platter holds 3.35 GB of data, with a density of 2.6 gigabit per square inch. The data transfer rate is about 13 MB per second. To put these numbers in perspective, observe that a disk access takes about 10 msecs, whereas accessing a main memory location typically takes less than 60 nanoseconds!

read is faulty for some reason, it is very unlikely that the checksum computed when the sector is read matches the checksum computed when the sector was written. The controller computes checksums, and if it detects an error, it tries to read the sector again. (Of course, it signals a failure if the sector is corrupted and read fails repeatedly.)

While direct access to any desired location in main memory takes approximately the same time, determining the time to access a location on disk is more complicated. The time to access a disk block has several components. **Seek time** is the time taken to move the disk heads to the track on which a desired block is located. As the size of a platter decreases, seek times also decrease, since we have to move a disk head a shorter distance. Typical platter diameters are 3.5 inches and 5.25 inches. **Rotational delay** is the waiting time for the desired block to rotate under the disk head; it is the time required for half a rotation all average and is usually less than seek time. **Transfer time** is the time to actually read or write the data in the block once the head is positioned, that is, the time for the disk to rotate over the block.

9.1.2 Performance Implications of Disk Structure

1. Data must be in mernory for the DBMS to operate on it.
2. The unit for data transfer between disk and main memory is a block; if a single item on a block is needed, the entire block is transferred. Reading or writing a disk block is called an I/O (for input/output) operation.
3. The time to read or write a block varies, depending on the location of the data:

$$access\ time = seek\ time + rotational\ delay + transfer\ time$$

These observations imply that the time taken for database operations is affected significantly by how data is stored on disks. The time for moving blocks to

or from disk usually dOlninates the time taken for database operations. To minimize this time, it is necessary to locate data records strategically on disk because of the geometry and mechanics of disks. **In** essence, if two records are frequently used together, we should place them close together. The 'closest' that two records can be on a disk is to be on the same block. **In** decreasing order of closeness, they could be on the same track, the same cylinder, or an adjacent cylinder.

Two records on the same block are obviously as close together as possible, because they are read or written as part of the same block. As the platter spins, other blocks on the track being read or written rotate under the active head. In current disk designs, all the data on a track can be read or written in one revolution. After a track is read or written, another disk head becomes active, and another track in the same cylinder is read or written. This process continues until all tracks in the current cylinder are read or written, and then the arm assembly moves (in or out) to an adjacent cylinder. Thus, we have a natural notion of 'closeness' for blocks, which we can extend to a notion of *next* and *previous* blocks.

Exploiting this notion of next by arranging records so they are read or written sequentially is very important in reducing the time spent in disk l/Os. Sequential access minimizes seek time and rotational delay and is much faster than random access. (This observation is reinforced and elaborated in Exercises 9.5 and 9.6, and the reader is urged to work through them.)

9.2 REDUNDANT ARRAYS OF INDEPENDENT DISKS

Disks are potential bottlenecks for system performance and storage system reliability. Even though disk performance has been improving continuously, microprocessor performance has advanced much more rapidly. The performance of microprocessors has improved at about 50 percent or more per year, but disk access times have improved at a rate of about 10 percent per year and disk transfer rates at a rate of about 20 percent per year. **In** addition, since disks contain mechanical elements, they have much higher failure rates than electronic parts of a computer system. **If** a disk fails, all the data stored on it is lost.

A **disk array** is an arrangement of several disks, organized to increase performance and improve reliability of the resulting storage system. Performance is increased through data striping. Data striping distributes data over several disks to give the impression of having a single large, very fast disk. Reliability is improved through **redundancy.** Instead of having a single copy of the data, redundant information is maintained. The redundant information is care-

fully organized so that, in case of a disk failure, it can be used to reconstruct the contents of the failed disk. Disk arrays that implement a combination of data striping and redundancy are called redundant arrays of independent disks, or in short, **RAID**.[1] Several RAID organizations, referred to as RAID levels, have been proposed. Each RAID level represents a different trade-off between reliability and performance.

In the remainder of this section, we first discuss data striping and redundancy and then introduce the RAID levels that have become industry standards.

9.2.1 Data Striping

A disk array gives the user the abstraction of having a single, very large disk. If the user issues an I/O request, we first identify the set of physical disk blocks that store the data requested. These disk blocks may reside on a single disk in the array or may be distributed over several disks in the array. Then the set of blocks is retrieved from the disk(s) involved. Thus, how we distribute the data over the disks in the array influences how many disks are involved when an I/O request is processed.

In data striping, the data is segmented into equal-size partitions distributed over multiple disks. The size of the partition is called the striping unit. The partitions are usually distributed using a round-robin algorithm: If the disk array consists of D disks, then partition i is written onto disk i mod D.

As an example, consider a striping unit of one bit. Since any D successive data bits are spread over all D data disks in the array, all I/O requests involve all disks in the array. Since the smallest unit of transfer from a disk is a block, each I/O request involves transfer of at least D blocks. Since we can read the D blocks from the D disks in parallel, the transfer rate of each request is D times the transfer rate of a single disk; each request uses the aggregated bandwidth of all disks in the array. But the disk access time of the array is basically the access time of a single disk, since all disk heads have to move for all requests. Therefore, for a disk array with a striping unit of a single bit, the number of requests per time unit that the array can process and the average response time for each individual request are similar to that of a single disk.

As another exarhple, consider a striping unit of a disk block. In this case, I/O requests of the size of a disk block are processed by one disk in the array. If many I/O requests of the size of a disk block are made, and the requested

[1] Historically, the *J* in RAID stood for inexpensive, as a large number of small disks was much more economical than a single very large disk. Today, such very large disks are not even manufactured—a sign of the impact of RAID.

Redundancy Schemes: Alternatives to the parity scheme include schemes based on Hamming codes and Reed-Solomon codes. In addition to recovery from single disk failures, Hamming codes can identify which disk failed. Reed-Solomon codes can recover from up to two simultaneous disk failures. A detailed discussion of these schemes is beyond the scope of our discussion here; the bibliography provides pointers for the interested reader.

blocks reside on different disks, we can process all requests in parallel and thus reduce the average response time of an I/O request. Since we distributed the striping partitions round-robin, large requests of the size of many contiguous blocks involve all disks. We can process the request by all disks in parallel and thus increase the transfer rate to the aggregated bandwidth of all D disks.

9.2.2 Redundancy

While having more disks increases storage system performance, it also lowers overall storage system reliability. Assume that the mean-time-to-failure (MTTF), of a single disk is 50,000 hours (about 5.7 years). Then, the MTTF of an array of 100 disks is only 50,000/100 = 500 hours or about 21 days, assuming that failures occur independently and the failure probability of a disk does not change over time. (Actually, disks have a higher failure probability early and late in their lifetimes. Early failures are often due to undetected manufacturing defects; late failures occur since the disk wears out. Failures do not occur independently either: consider a fire in the building, an earthquake, or purchase of a set of disks that come from a 'bad' manufacturing batch.)

Reliability of a disk array can be increased by storing redundant information. If a disk fails, the redundant information is used to reconstruct the data on the failed disk. Redundancy can immensely increase the MTTF of a disk array. When incorporating redundancy into a disk array design, we have to make two choices. First, we have to decide where to store the redundant information. We can either store the redundant information on a small number of check disks or distribute the redundant information uniformly over all disks.

The second choice we have to make is how to compute the redundant information. Most disk arrays store parity information: In the parity scheme, an extra check disk contains information that can be used to recover from failure of anyone disk in the array. Assume that we have a disk array with D disks and consider the first bit on each data disk. Suppose that i of the D data bits are 1. The first bit on the check disk is set to 1 if i is odd; otherwise, it is set to

0. This bit on the check disk is called the parity of the data bits. The check disk contains parity information for each set of corresponding D data bits.

To recover the value of the first bit of a failed disk we first count the number of bits that are 1 on the D - 1 nonfailed disks; let this number be *j*. If *j* is odd and the parity bit is 1, or if *j* is even and the parity bit is 0, then the value of the bit on the failed disk must have been 0. Otherwise, the value of the bit on the failed disk must have been 1. Thus, with parity we can recover from failure of anyone disk. Reconstruction of the lost information involves reading all data disks and the check disk.

For example, with an additional 10 disks with redundant information, the MTTF of our example storage system with 100 data disks can be increased to more than 250 years! "What is more important, a large MTTF implies a small failure probability during the actual usage time of the storage system, which is usually much smaller than the reported lifetime or the MTTF. (Who actually uses lO-year-old disks?)

In a RAID system, the disk array is partitioned into reliability groups, where a reliability group consists of a set of *data disks* and a set of *check disks.* A common *7'cdundancy scheme* (see box) is applied to each group. The number of check disks depends on the RAID level chosen. In the remainder of this section, we assume for ease of explanation that there is only one reliability group. The reader should keep in mind that actual RAID implementations consist of several reliability groups, and the number of groups plays a role in the overall reliability of the resulting storage system.

9.2.3 Levels of Redundancy

Throughout the discussion of the different RAID levels, we consider sample data that would just fit on four disks. That is, with no RAID technology our storage system would consist of exactly four data disks. Depending on the RAID level chosen, the number of additional disks varies from zero to four.

Level 0: Nonredundant

A RAID Level 0 system uses data striping to increase the maximum bandwidth available. No redundant information is maintained. While being the solution with the lowest cost, reliability is a problem, since the MTTF decreases linearly with the number of disk drives in the array. RAID Level 0 has the best write performance of all RAID levels, because absence of redundant information implies that no redundant information needs to be updated! Interestingly, RAID Level 0 docs not have the best read perfonnancc of all RAID levels, since sys-

tems with redundancy have a choice of scheduling disk accesses, as explained in the next section.

In our example, the RAID Level a solution consists of only four data disks. Independent of the number of data disks, the effective space utilization for a RAID Level a system is always 100 percent.

Levell: Mirrored

A RAID Level 1 system is the most expensive solution. Instead of having one copy of the data, two identical copies of the data on two different disks are lnaintained. This type of redundancy is often called mirroring. Every write of a disk block involves a write on both disks. These writes may not be performed simultaneously, since a global system failure (e.g., due to a power outage) could occur while writing the blocks and then leave both copies in an inconsistent state. Therefore, we always write a block on one disk first and then write the other copy on the mirror disk. Since two copies of each block exist on different disks, we can distribute reads between the two disks and allow *parallel reads* of different disk blocks that conceptually reside on the same disk. A read of a block can be scheduled to the disk that has the smaller expected access time. RAID Level 1 does not stripe the data over different disks, so the transfer rate for a single request is comparable to the transfer rate of a single disk.

In our example, we need four data and four check disks with mirrored data for a RAID Levell implementation. The effective space utilization is 50 percent, independent of the number of data disks.

Level 0+1: Striping and Mirroring

RAID Level 0+l---sometimes also referred to as *RAID Level 10*--combines striping and mirroring. As in RAID Level 1. read requests of the size of a disk block can be scheduled both to a disk and its mirror image. In addition, read requests of the size of several contiguous blocks benefit froll1 the aggregated bandwidth of all disks. Thc cost for writes is analogous to RAID LevelL

As in RAID Level 1, our example with four data disks requires four check disks and the effective space utilization is always 50 percent.

Level 2: **Error-Correcting** Codes

In RAID Level 2, the striping unit is a single bit. The redundancy scheme used is Hamming code. In our example with four data disks, only three check disks

are needed. In general, the number of check disks grows logarithmically with the number of data disks.

Striping at the bit level has the implication that in a disk array with D data disks, the smallest unit of transfer for a read is a set of D blocks. Therefore, Level 2 is good for workloads with many large requests, since for each request, the aggregated bandwidth of all data disks is used. But RAID Level 2 is bad for small requests of the size of an individual block for the same reason. (See the example in Section 9.2.1.) A write of a block involves reading D blocks into main memory, modifying $D + C$ blocks, and writing $D + C$ blocks to disk, where C is the number of check disks. This sequence of steps is called a *read-modify-write* cycle.

For a RAID Level 2 implementation with four data disks, three check disks are needed. In our example, the effective space utilization is about 57 percent. The effective space utilization increases with the number of data disks. For example, in a setup with 10 data disks, four check disks are needed and the effective space utilization is 71 percent. In a setup with 25 data disks, five check disks are required and the effective space utilization grows to 83 percent.

Level 3: **Bit-Interleaved Parity**

While the redundancy schema used in RAID Level 2 improves in terms of cost over RAID Level 1, it keeps more redundant information than is necessary. Hamming code, as used in RAID Level 2, has the advantage of being able to identify which disk has failed. But disk controllers can easily detect which disk has failed. Therefore, the check disks do not need to contain information to identify the failed disk. Information to recover the lost data is sufficient. Instead of using several disks to store Hamming code, RAID Level 3 has a single check disk with parity information. Thus, the reliability overhead for RAID Level 3 is a single disk, the lowest overhead possible.

The performance characteristics of RAID Levels 2 and 3 are very similar. RAID Level 3 can also process only one I/O at a time, the minimum transfer unit is D blocks, and a write requires a read-modify-write cycle.

Level 4: **Block-Interleaved Parity**

RAID Level 4 has a striping unit of a disk block, instead of a single bit as in RAID Level 3. Block-level striping has the advantage that read requests of the size of a disk block can be served entirely by the disk where the requested block resides. Large read requests of several disk blocks can still utilize the aggregated bandwidth of the D disks.

The \vrite of a single block still requires a read-modify-write cycle, but only one data disk and the check disk are involved. The parity on the check disk can be updated without reading all D disk blocks, because the new parity can be obtained by noticing the differences between the old data block and the new data block and then applying the difference to the parity block on the check disk:

NewParity = (OldData XOR NewData) XOR OldParity

The read-modify-write cycle involves reading of the old data block and the old parity block, modifying the two blocks, and writing them back to disk, resulting in four disk accesses per write. Since the check disk is involved in each write, it can easily become the bottleneck.

RAID Level 3 and 4 configurations with four data disks require just a single check disk. In our example, the effective space utilization is 80 percent. The effective space utilization increases with the number of data disks, since always only one check disk is necessary.

Level 5: **Block-Interleaved Distributed Parity**

RAID Level 5 improves on Level 4 by distributing the parity blocks uniformly over all disks, instead of storing them on a single check disk. This distribution has two advantages. First, several write requests could be processed in parallel, since the bottleneck of a unique check disk has been eliminated. Second, read requests have a higher level of parallelism. Since the data is distributed over all disks, read requests involve all disks, whereas in systems with a dedicated check disk the check disk never participates in reads.

A RAID Level 5 system has the best performance of all RAID levels with redundancy for small and large read ancllarge write requests. Small writes still require a read-modify-write cycle and are thus less efficient than in RAID Level 1.

In our example, the corresponding RAID Level 5 system has five disks overall and thus the effective spa,ce utilization is the same as in RAID Levels 3 and 4.

Level 6: P+Q Redundancy

The motivation for RAID Level 6 is the observation that recovery from failure of a single disk is not sufficient in very large disk arrays. First, in large disk arrays, a second disk lllight fail before replacement of an already failed disk

could take place. In addition, the probability of a disk failure during recovery of a failed disk is not negligible.

A RAID Level 6 system uses Reed-Solomon codes to be able to recover from up to two simultaneous disk failures. RAID Level 6 requires (conceptually) two check disks, but it also uniformly distributes redundant information at the block level as in RAID Level 5. Thus. the performance characteristics for small and large read requests and for large write requests are analogous to RAID Level 5. For small writes, the read-modify-write procedure involves six instead of four disks as compared to RAID Level 5, since two blocks with redundant information need to be updated.

For a RAID Level 6 system with storage capacity equal to four data disks, six disks are required. In our example, the effective space utilization is 66 percent.

9.2.4 Choice of RAID Levels

If data loss is not an issue, RAID Level 0 improves overall system performance at the lowest cost. RAID Level 0+1 is superior to RAID Level 1. The main application areas for RAID Level 0+1 systems are small storage subsystems where the cost of mirroring is moderate. Sometimes, RAID Level 0+1 is used for applications that have a high percentage of writes in their workload, since RAID Level 0+1 provides the best write performance. RAID Levels 2 and 4 are always inferior to RAID Levels 3 and 5, respectively. RAID Level 3 is appropriate for workloads consisting mainly of large transfer requests of several contiguous blocks. The performance of a RAID Level 3 system is bad for workloads with many small requests of a single disk block. RAID Level 5 is a good general-purpose solution. It provides high performance for large as well as small requests. RAID Level 6 is appropriate if a higher level of reliability is required.

9.3 DISK SPACE MANAGEMENT

The lowest level of software in the DBMS architecture discussed in Section 1.8, called the disk space manager, manages space on disk. Abstractly, the disk space manager supports the concept of a page as a unit of data and provides commands to allocate or deallocate a page and read or write a page. The size of a page is chosen to be the size of a disk block and pages are stored as disk blocks so that reading or writing a page can be done in one disk I/O.

It is often useful to allocate a sequence of pages as a *contiguous* sequence of blocks to hold data frequently accessed in sequential order. This capability is essential for exploiting the advantages of sequentially accessing disk blocks,

which we discussed earlier in this chapter. Such a capability, if desired, must be provided by the disk space manager to higher-level layers of the DBMS.

The disk space manager hides details of the underlying hardware (and possibly the operating system) and allows higher levels of the software to think of the data as a collection of pages.

9.3.1 Keeping Track of Free Blocks

A database grows and shrinks as records are inserted and deleted over time. The disk space manager keeps track of which disk blocks are in usc, in addition to keeping track of which pages are on which disk blocks. Although it is likely that blocks are initially allocated sequentially on disk, subsequent allocations and deallocations could in general create 'holes.'

One way to keep track of block usage is to maintain a list of free blocks. As blocks are deallocated (by the higher-level software that requests and uses these blocks), we can add them to the free list for future use. A pointer to the first block on the free block list is stored in a known location on disk.

A second way is to maintain a bitmap with one bit for each disk block, which indicates whether a block is in use or not. A bitmap also allows very fast identification and allocation of contiguous areas on disk. This is difficult to accomplish with a linked list approach.

9.3.2 Using OS File Systems to Manage Disk Space

Operating systems also manage space on disk. Typically, an operating system supports the abstraction of a *file as a sequence of bytes.* The as manages space on the disk and translates requests, such as "Read byte i of file *f*," into corresponding low-level instructions: "Read block m of track *t* of cylinder c of disk *d.*" A database disk space manager could he built using OS files. For example, the entire database could reside in one or more as files for which a number of blocks are allocated (by the aS) and initialized. The disk space manager is then responsible for managing the space in these OS files.

Many database systems do not rely on the as file system and instead do their own disk management, either from scratch or by extending as facilities. The reasons are practical as well as technical One practical reason is that a DBMS vendor who wishes to support several as platfonns cannot assume features specific to any OS, for portability, and would therefore try to make the DBMS code as self-contained as possible. A technical reason is that on a :32-bit systern, the la.rgest file size is 4 GB. whereas a DBMS may want to access a single file

larger than that. A related problem is that typical as files cannot span disk devices, which is often desirable or even necessary in a DBMS. Additional technical reasons why a DBMS does not rely on the as file system are outlined in Section 9.4.2.

9.4 BUFFER MANAGER

To understand the role of the buffer manager, consider a simple example. Suppose that the database contains 1 million pages, but only 1000 pages of main memory are available for holding data. Consider a query that requires a scan of the entire file. Because all the data cannot be brought into main memory at one time, the DBMS must bring pages into main memory as they are needed and, in the process, decide what existing page in main memory to replace to make space for the new page. The policy used to decide which page to replace is called the replacement policy.

In terms of the DBMS architecture presented in Section 1.8, the buffer manager is the software layer responsible for bringing pages from disk to main memory as needed. The buffer manager manages the available main memory by partitioning it into a collection of pages, which we collectively refer to as the buffer pool. The main memory pages in the buffer pool are called frames; it is convenient to think of them as slots that can hold a page (which usually resides on disk or other secondary storage media).

Higher levels of the DBMS code can be written without worrying about whether data pages are in memory or not; they ask the buffer manager for the page, and it is brought into a frame in the buffer pool if it is not already there. Of course, the higher-level code that requests a page must also release the page when it is no longer needed, by informing the buffer manager, so that the frame containing the page can be reused. The higher-level code must also inform the buffer manager if it modifies the requested page; the buffer manager then makes sure that the change is propagated to the copy of the page on disk. Buffer management is illustrated in Figure 9.3.

In addition to the buffer pool itself, the buffer manager maintains some bookkeeping information and two variables for each frame in the pool: *pirLcount* and *dirty*. The number of times that the page currently in a given frame has been requested but not released—the number of current users of the page--is recorded in the *pin_count* variable for that frame. The Boolean variable *dirty* indicates whether the page has been modified since it was brought into the buffer pool from disk.

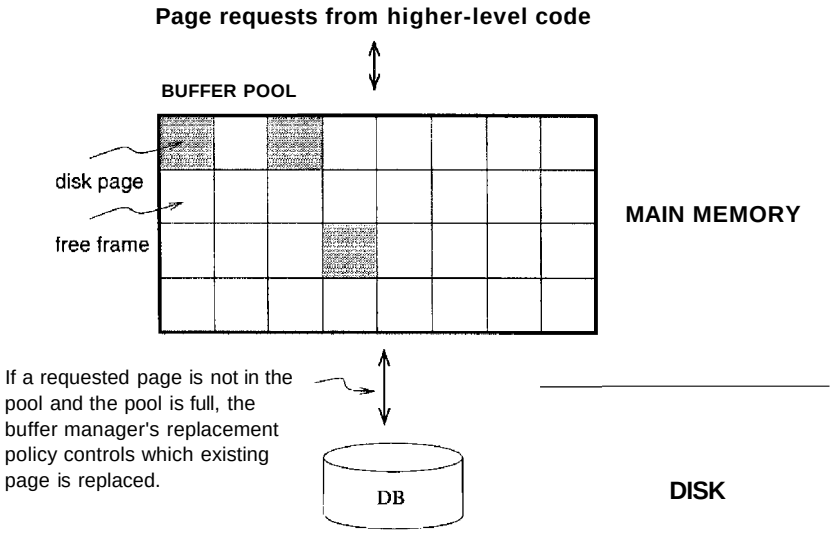


Figure 9.3 The Buffer Pool

Initially, the *pin_count* for every frame is set to 0, and the *dirty* bits are turned off. When a page is requested the buffer manager does the following:

1. Checks the buffer pool to see if some frame contains the requested page and, if so, increments the *pin_count* of that frame. If the page is not in the pool, the buffer manager brings it in as follows:
   (a) Chooses a frame for replacement, using the replacement policy, and increments its *pin_count.*
   (b) If the *dirty* bit for the replacement frame is on, writes the page it contains to disk (that is, the disk copy of the page is overwritten with the contents of the frame).
   (c) Reads the requested page into the replacement frame.
2. Returns the (main memory) address of the frame containing the requested page to the requestor.

Incrementing *pin_count* is often called **pinning** the requested page in its frame. When the code that calls the buffer manager and requests the page subsequently calls the buffer manager and releases the page, the *pin_count* of the frame containing the requested page is decremented. This is called **unpinning** the page. If the requestor has modified the page, it also informs the buffer manager of this at the time that it unpins the page, and the *dirty* bit for the frame is set.

The buffer manager will not read another page into a frame until its *pin_count* becomes 0, that is, until all requestors of the page have unpinned it.

If a requested page is not in the buffer pool and a free frame is not available in the buffer pool, a frame with *pin_count* 0 is chosen for replacement. If there are many such frames, a frame is chosen according to the buffer manager's replacement policy. We discuss various replacement policies in Section 9.4.1.

When a page is eventually chosen for replacement, if the *dirty* bit is not set, it means that the page has not been modified since being brought into main memory. Hence, there is no need to write the page back to disk; the copy on disk is identical to the copy in the frame, and the frame can simply be overwritten by the newly requested page. Otherwise, the modifications to the page must be propagated to the copy on disk. (The crash recovery protocol may impose further restrictions, as we saw in Section 1.7. For example, in the Write-Ahead Log (WAL) protocol, special log records are used to describe the changes made to a page. The log records pertaining to the page to be replaced may well be in the buffer; if so, the protocol requires that they be written to disk *before* the page is written to disk.)

If no page in the buffer pool has *pin_count* 0 and a page that is not in the pool is requested, the buffer manager must wait until some page is released before responding to the page request. In practice, the transaction requesting the page may simply be aborted in this situation! So pages should be released—by the code that calls the buffer manager to request the page- as soon as possible.

A good question to ask at this point is, "What if a page is requested by several different transactions?" That is, what if the page is requested by programs executing independently on behalf of different users? Such programs could make conflicting changes to the page. The locking protocol (enforced by higher-level DBMS code, in particular the transaction manager) ensures that each transaction obtains a shared or exclusive lock before requesting a page to read or rnodify. Two different transactions cannot hold an exclusive lock on the same page at the same time; this is how conflicting changes are prevented. The buffer rnanager simply assumes tha.t the appropriate lock has been obtained before a page is requested.

9.4.1 Buffer Replacement Policies

The policy used to choose an unpinned page for replacement can affect the time taken for database operations considerably. Of the man,Y alternative policies, each is suitable in different situations.

The best-known replacement policy is least recently used (LRU). This can be implemented in the buffer manager using a queue of pointers to frames with *pin_count* 0. A frame is added to the end of the queue when it becomes a candidate for replacement (that is, when the *pin_count* goes to 0). The page chosen for replacement is the one in the frame at the head of the queue.

A variant of LRU, called clock replacement, has similar behavior but less overhead. The idea is to choose a page for replacement using a *current* variable that takes on values 1 through *N*, where *N* is the number of buffer frames, in circular order. We can think of the frames being arranged in a circle, like a clock's face, and *current* as a clock hand moving across the face. To approximate LRU behavior, each frame also has an associated *referenced* bit, which is turned on when the page *pin_count* goes to 0.

The *current* frame is considered for replacement. If the frame is not chosen for replacement, *current* is incremented and the next frame is considered; this process is repeated until some frame is chosen. If the *current* frame has *pin_count* greater than 0, then it is not a candidate for replacement and *current* is incremented. If the *current* frame has the *referenced* bit turned on, the clock algorithm turns the *referenced* bit off and increments *current*—this way, a recently referenced page is less likely to be replaced. If the *current* frame has *pin_count* 0 and its *referenced* bit is off, then the page in it is chosen for replacement. If all frames are pinned in some sweep of the clock hand (that is, the value of *current* is incremented until it repeats), this means that no page in the buffer pool is a replacement candidate.

The LRU and clock policies are not always the best replacement strategies for a database system, particularly if many user requests require sequential scans of the data. Consider the following illustrative situation. Suppose the buffer pool has 10 frames, and the file to be scanned has 10 or fewer pages. Assuming, for simplicity, that there are no competing requests for pages, only the first scan of the file does any I/O. Page requests in subsequent scans always find the desired page in the buffer pool. On the other hand, suppose that the file to be scanned has 11 pages (which is one more than the number of available pages in the buffer pool). Using LRU, every scan of the file will result in reading every page of the file! In this situation, called sequential flooding, LRU is the *worst* possible replacement strategy.

Other replacement policies include first in first out (FIFO) and most recently used (MRU), which also entail overhead similar to LRU, and random, arnong others. The details of these policies should be evident from their names and the preceding discussion of LRU and clock.

Buffer Management in Practice: IBM DB2 and Sybase ASE allow buffers to be partitioned into named pools. Each database, table, or index can be bound to one of these pools. Each pool can be configured to use either LRU or clock replacement in ASE; DB2 uses a variant of clock replacement, with the initial clock value based on the nature of the page (e.g., index non-leaves get a higher starting clock value, which delays their replacement). Interestingly, a buffer pool client in DB2 can explicitly indicate that it *hates* a page, making the page the next choice for replacement. As a special case, DB2 applies MRU for the pages fetched in some utility operations (e.g., RUNSTATS), and DB2 V6 also supports FIFO. Informix and Oracle 7 both maintain a single global buffer pool using LRU; Microsoft SQL Server has a single pool using clock replacement. In Oracle 8, tables can be bound to one of two pools; one has high priority, and the system attempts to keep pages in this pool in memory.

Beyond setting a maximum number of pins for a given transaction, there are typically no features for controlling buffer pool usage on a per-transaction basis. Microsoft SQL Server, however, supports a reservation of buffer pages by queries that require large amounts of memory (e.g., queries involving sorting or hashing).

9.4.2 Buffer Management in DBMS versus OS

Obvious similarities exist between virtual memory in operating systems and buffer management in database management systems. In both cases, the goal is to provide access to more data than will fit in main memory, and the basic idea is to bring in pages from disk to main memory as needed, replacing pages no longer needed in main memory. Why can't we build a DBMS using the virtual memory capability of an OS? A DBMS can often predict the order in which pages will be accessed, or page reference patterns, much more accurately than is typical in an as environment, and it is desirable to utilize this property. Further, a DBMS needs more control over when a page is written to disk than an as typically provides.

A DBMS can often predict reference patterns because most page references are generated by higher-level operations (such as sequential scans or particular implementations of various relational algebra opera.tors) with a known pattern of page accesses. This ability to predict reference patterns allows for a better choice of pages to replace and makes the idea of specialized buffer replacement policies more attractive in the DBMS environment.

Even more important, being able to predict reference patterns enables the use of a simple and very effective strategy called prefetching of pages. The

Prefetching: IBM DB2 supports both sequential alld list prefeteh (prefetching a list of pages). In general, the prefeteh size is 32 4KB pages, but this can be set by the user. For some sequential type datahaseutilities (e.g., COPY, RUNSTATS), DB2 prefetches up to 64 4KB pages. For a smaller buffer pool (i.e., less than 1000 buffers), the prefetch quantity is adjusted downward to 16 or 8 pages. The prefetch size can be configured by the user; for certain environments, it may be best to prefetch 1000 pages at a time! Sybase ASE supports asynchronous prefetching of up to 256 pages, and uses this capability to reduce latency during indexed access to a table in a range scan. Oracle 8 uses prefetching for sequential scan, retrieving large objects, and certain index scans. Microsoft SQL Server supports prefetching for sequential scan and for scans along the leaf level ofa B+ tree index, and the prefetch size can be adjusted as a scan progresses. SQL Server also uses asynchronous prefetching extensively. Informix supports prefetching with a user-defined prefetch size.

buffer manager can anticipate the next several page requests and fetch the corresponding pages into memory *before* the pages are requested. This strategy has two benefits. First, the pages are available in the buffer pool when they are requested. Second, reading in a contiguous block of pages is much faster than reading the same pages at different times in response to distinct requests. (Review the discussion of disk geometry to appreciate why this is so.) If the pages to be prcfetched are not contiguous, recognizing that several pages need to be fetched can nonetheless lead to faster I/O because an order of retrieval can be chosen for these pages that minimizes seek times and rotational delays.

Incidentally, note that the I/O can typically be done concurrently with CPU computation. Once the prefetch request is issued to the disk, the disk is responsible for reading the requested pages into memory pages and the CPU can continue to do other work.

A DBMS also requires the ability to explicitly *force* a page to disk, that is, to ensure that the copy of the page on disk is updated with the copy in memory. As a related point, a DBMS must be able to ensure that certain pages in the buffer pool are written to disk *before* certain other pages to implement the WAL protocol for crash recovery, as we saw in Section 1.7. Virtual memory implementations in operating systems cannot be relied on to provide such control over when pages are written to disk; the OS command to write a page to disk may be implemented by essentially recording the write request and deferring the actual modification of the disk copy. If the systern crashes in the interim, the effects can be catastrophic for a DBMS. (Crash recovery is discllssed further in Chapter 18.)

Indexes as Files: In Chapter 8, we presented indexes as a way of organizing data records for efficient search. From an implementation standpoint, indexes are just another kind of file, containing records that dil'ect traffic on requests for data records. For example, a tree index is a collection of records organized into one page per node in the tree. It is convenient to actually think of a tree index as *two* files, because it contains two kinds of records: (1) a file of *index entries*, which are records with fields for the index's search key, and fields pointing to a child node, and (2) a file of *data entries*, whose structure depends on the choice of data entry alternative.

9.5 FILES OF RECORDS

We now turn our attention from the way pages are stored on disk and brought into main memory to the way pages are used to store records and organized into logical collections or *files*. Higher levels of the DBMS code treat a page as effectively being a collection of records, ignoring the representation and storage details. In fact, the concept of a collection of records is not limited to the contents of a single page; a file can span several pages. In this section, we consider how a collection of pages can be organized as a file. We discuss how the space on a page can be organized to store a collection of records in Sections 9.6 and 9.7.

9.5.1 Implementing Heap Files

The data in the pages of a heap file is not ordered in any way, and the only guarantee is that one can retrieve all records in the file by repeated requests for the next record. Every record in the file has a unique rid, and every page in a file is of the same size.

Supported operations on a heap file include *CTeatc* and *destroy* files, *insert* a record, *delete* a record with a given rid, *get* a record with a given rid, and *scan* all records in the file. To get or delete a record with a given rid, note that we must be able to find the id of the page containing the record, given the id of the record.

We must keep track of the pages in each heap file to support scans, and we must keep track of pages that contain free space to implement insertion efficiently. \Ve discuss two alternative ways to rnaintain this information. In each of these alternatives, pages must hold two pointers (which are page ids) for file-level bookkeeping in addition to the data.

Linked List of **Pages**

One possibility is to maintain a heap file as a doubly linked list of pages. The DBMS can remember where the first page is located by maintaining a table containing pairs of $\langle heap\_file\_name, page\_Laddr \rangle$ in a known location on disk. We call the first page of the file the *header page.*

An important task is to maintain information about empty slots created by deleting a record from the heap file. This task has two distinct parts: how to keep track of free space within a page and how to keep track of pages that have some free space. We consider the first part in Section 9.6. The second part can be addressed by maintaining a doubly linked list of pages with free space and a doubly linked list of full pages; together, these lists contain *all* pages in the heap file. This organization is illustrated in Figure 9.4; note that each pointer is really a page id.

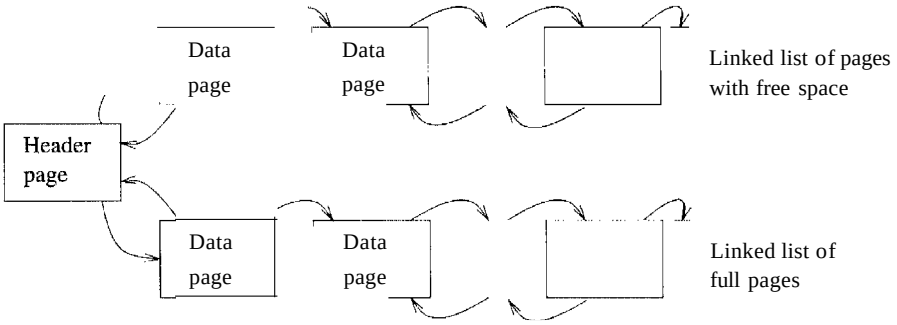


Figure 9.4 Heap File Organization with a Linked List

If a new page is required, it is obtained by making a request to the disk space manager and then added to the list of pages in the file (probably as a page with free space, because it is unlikely that the new record will take up all the space on the page). **If** a page is to be deleted from the heap file, it is removed from the list and the disk space Inanager is told to deallocate it. (Note that the scheme can easily be generalized to allocate or deallocate a sequence of several pages and maintain a doubly linked list of these page sequences.)

One disadvantage of this scheIue is that virtually all pages in a file will be on the free list if records are of variable length, because it is likely that every page has at least a few free bytes. To insert a typical record, we must retrieve and exaInine several pages on the free list before we find one with enough free space. The directory-based heap file organization that we discuss next addresses this problem.

Directory of Pages

An alternative to a linked list of pages is to maintain a **directory of** pages. The DBMS must remember where the first directory page of each heap file is located. The directory is itself a collection of pages and is shown as a linked list in Figure 9.5. (Other organizations are possible for the directory itself, of course.)

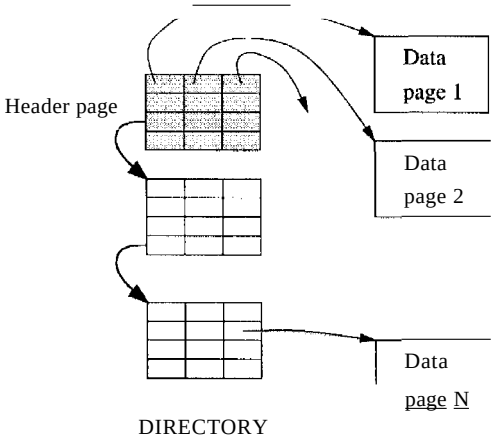


Figure 9.5 Heap File Organization with a Directory

Each directory entry identifies a page (or a sequence of pages) in the heap file. As the heap file grows or shrinks, the number of entries in the directory-and possibly the number of pages in the directory itself--grows or shrinks correspondingly. Note that since each directory entry is quite small in comparison to a typical page, the size of the directory is likely to be very small in comparison to the size of the heap file.

Free space can be managed by maintaining a bit per entry, indicating whether the corresponding page has any free space, or a count per entry, indicating the amount of free space on the page. If the file contains variable-length records, we can examine the free space count for an entry to determine if the record fits on the page pointed to by the entry. Since several entries fit on a directory page, we can efficiently search for a data page with enough space to hold a record to be inserted.

9.6 PAGE FORMATS

The page abstraction is appropriate when dealing with I/O issues, but higher levels of the DBMS see data as a collection of records. In this section, we

Rids in COInmercial Systems: IBM DB2, Informix, Microsoft SQL Server, Oracle 8, and Sybase ASE all implement record ids as a page id and slot number. Sybase ASE uses the following page organization, which is typical: Pages contain a header followed by the rows and a slot array. The header contains the page identity, its allocation state, page free space state, and a timestamp. The slot array is simply a mapping of slot number to page offset.

Oracle 8 and SQL Server use logical record ids rather than page id and slot number in one special case: If a table has a clustered index, then records in the table are identified using the key value for the clustered index. This has the advantage that secondary indexes need not be reorganized if records are moved across pages.

consider how a collection of records can be arranged on a page. We can think of a page as a collection of slots, each of which contains a record. A record is identified by using the pair *(page id, slot number);* this is the record id (rid). (We remark that an alternative way to identify records is to assign each record a unique integer as its rid and maintain a table that lists the page and slot of the corresponding record for each rid. Due to the overhead of maintaining this table, the approach of using *(page id, slot number)* as an rid is more common.)

We now consider some alternative approaches to managing slots on a page. The main considerations are how these approaches support operations such as searching, inserting, or deleting records on a page.

9.6.1 Fixed-Length **Records**

If all records on the page are guaranteed to be of the same length, record slots arc uniform and can be arranged consecutively within a page. At any instant, some slots are occupied by records and others are unoccupied. When a record is inserted into the page, we must locate an empty slot and place the record there. The main issues are how we keep track of empty slots and how we locate all records on a page. The alternatives hinge on how we handle the deletion of a record.

The first alternative is to store records in the first N slots (where N is the number of records on the page); whenever a record is deleted, we move the last record on the page into the vacated slot. This format allows us to locate the ith record on a page by a simple offset calculation, and all empty slots appear together at the end of the page. However, this approach docs not work if there

are external references to the record that is moved (because the rid contains the slot number, which is now changed).

The second alternative is to handle deletions by using an array of bits, one per slot, to keep track of free slot information. Locating records on the page requires scanning the bit array to find slots whose bit is on; when a record is deleted, its bit is turned off. The two alternatives for storing fixed-length records are illustrated in Figure 9.6. Note that in addition to the information about records on the page, a page usually contains additional file-level information (e.g., the id of the next page in the file). The figure does not show this additional information.

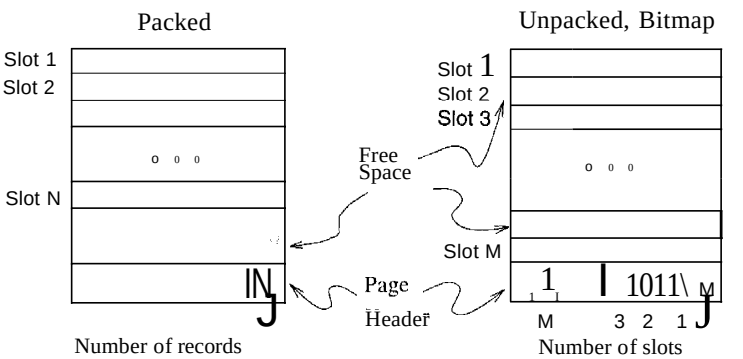


Figure 9.6 Alternative Page Organizations for Fixed-Length Recorcls

The *slotted page* organization described for variable-length records in Section 9.6.2 can also be used for fixed-length records. It becomes attractive if we need to move records around on a page for reasons other than keeping track of space freed by deletions. A typical example is that we want to keep the records on a page sorted (according to the value in some field).

9.6.2 Variable-Length Records

If records are of variable length, then we cannot divide the page into a fixed collection of slots. The problem is that, when a new record is to be inserted, we have to find an empty slot of just the right length----if we use a slot that is too big, we waste space, ancl obviously we cannot use a slot that is smaller than the record length. Therefore, when a record is inserted, we must allocate just the right amount of space for it, and when a record is deleted, we must move records to fill the hole created by the deletion, to ensure that all the free space on the page is contiguous. Therefore, the ability to move records on a page becomes very important.

The most flexible organization for variable-length records is to maintain a directory of slots for each page, with a *(record offset, record length)* pair per slot. The first component *(record offset)* is a 'pointer' to the record, as shown in Figure 9.7; it is the offset in bytes from the start of the data area on the page to the start of the record, Deletion is readily accomplished by setting the record offset to -1. Records can be moved around on the page because the rid, which is the page number and slot number (that is, position in the directory), does not change when the record is moved; only the record offset stored in the slot changes.

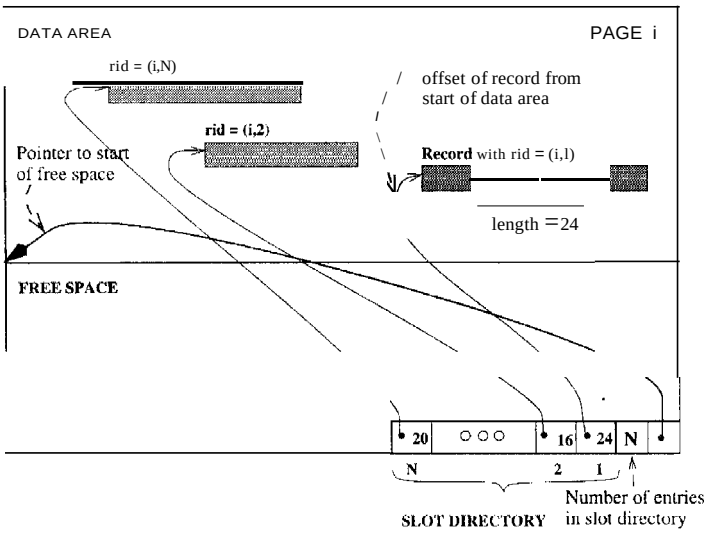


Figure 9.7 Page Organization for Variable-Length R.ecords

The space available for new records must be managed carefully because the page is not preformatted into slots. One way to manage free space is to maintain a pointer (that is, offset from the start of the data area on the page) that indicates the start of the free space area. When a new record is too large to fit into the remaining free space, we have to move records on the page to reclairn the space freed by records deleted earlier. The idea is to ensure that, after reorganization, all records appear in contiguous order, followed by the available free space.

A subtle point to be noted is that the slot for a deleted record cannot always be removed from the slot directory, because slot numbers are used to identify records---by deleting a slot, we change (decrement) the slot number of subsequent slots in the slot directory, and thereby change the rid of records pointed to by subsequent slots. The only way to remove slots from the slot directory is to remove the last slot if the record that it points to is deleted. However, when

a record is inserted, the slot directory should be scanned for an element that currently does not point to any record, and this slot should be used for the new record. A new slot is added to the slot directory only if all existing slots point to records. If inserts are much more common than deletes (as is typically the case), the number of entries in the slot directory is likely to be very close to the actual number of records on the page.

This organization is also useful for fixed-length records if we need to move them around frequently; for example, when we want to maintain them in some sorted order. Indeed, when all records are the same length, instead of storing this common length information in the slot for each record, we can store it once in the system catalog.

In some special situations (e.g., the internal pages of a B+ tree, which we discuss in Chapter 10), we lIlay not care about changing the rid of a record. In this case, the slot directory can be compacted after every record deletion; this strategy guarantees that the number of entries in the slot directory is the same as the number of records on the page. If we do not care about modifying rids, we can also sort records on a page in an efficient manner by simply moving slot entries rather than actual records, which are likely to be much larger than slot entries.

A simple variation on the slotted organization is to maintain only record offsets in the slots. For variable-length records, the length is then stored with the record (say, in the first bytes). This variation makes the slot directory structure for pages with fixed-length records the sallle as for pages with variable-length records.

9.7 RECORD FORMATS

In this section, we discuss how to organize fields within a record. While choosing a way to organize the fields of a record, we must take into account whether the fields of the record are of fixed or variable length and consider the cost of various operations on the record, including retrieval and modification of fields.

Before discussing record fonnats, we note that in addition to storing individual records, information conllnon to all records of a given record type (such as the number of fields and field types) is stored in the system catalog, which can be thought of as a description of the contents of a database, maintained by the DBMS (Section 12.1). This avoids repeated storage of the same information with each record of a given type.

Record Formats **in** Commercial Aystems: In IBM DB2, fixed-length fields are at fixed offsets from the beginning of the record. Variable-length fields have offset and length in the fixed offset part of the record, and the fields themselves follow the fixed-length part of the record. Informix, Microsoft SQL Server, and Sybase ASE use the same organization with minor variations. In Oracle 8, records are structured as if all fields are potentially of variable length; a record is a sequence of length-data pairs, with a special length value used to denote a *null* value.

9.7.1 Fixed-Length Records

In a fixed-length record, each field has a fixed length (that is, the value in this field is of the same length in all records), and the number of fields is also fixed. The fields of such a record can be stored consecutively, and, given the address of the record, the address of a particular field can be calculated using information about the lengths of preceding fields, which is available in the system catalog. This record organization is illustrated in Figure 9.8.

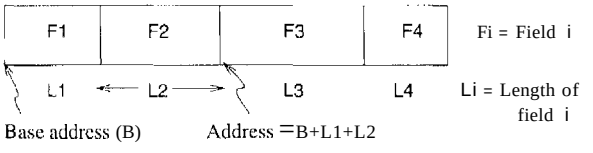


Figure 9.8 Organization of Records with Fixed-Length Fields

9.7.2 Variable-Length Records

In the relational model, every record in a relation contains the same number of fields. **If** the number of fields is fixed, a record is of variable length only because some of its fields are of variable length.

One possible orga,nizatioll is to store fields consecutively, separated by delimiters (which are special characters that do not appear in the data itself). This organization requires a scan of the record to locate a desired field.

An alternative is to reserve some space at the beginning of a record for use as an array of integer offsets—the ith integer in this array is the starting address of the ith field value relative to the start of the record. Note that we also store an offset to the end of the record; this offset is needed to recognize where the last field ends. Both alternatives are illustrated in Figure 9.9.

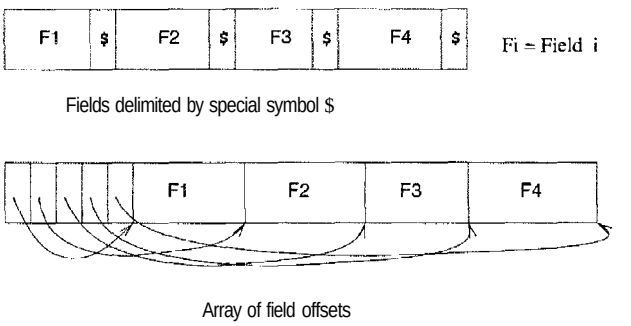


Figure 9.9 Alternative Record Organizations for Variable-Length Fields

The second approach is typically superior. For the overhead of the offset array, we get direct access to any field. We also get a clean way to deal with **null** values. A *null* value is a special value used to denote that the value for a field is unavailable or inapplicable. If a field contains a *null* value, the pointer to the end of the field is set to be the same as the pointer to the beginning of the field. That is, no space is used for representing the *null* value, and a comparison of the pointers to the beginning and the end of the field is used to determine that the value in the field is *null*.

Variable-length record formats can obviously be used to store fixed-length records as well; sometimes, the extra overhead is justified by the added flexibility, because issues such as supporting *n'ull* values and adding fields to a recorcl type arise with fixed-length records as well.

I-laving variable-length fields in a record can raise some subtle issues, especially when a record is modified.

- Modifying a field may cause it to grow, which requires us to shift all subsequent fields to make space for the modification in all three record formats just presentcel.
- A modified record may no longer fit into the space remaining on its page. If so, it may have to be moved to another page. If riels, which are used to 'point' to a record, include the page number (see Section 9.6), moving a record to 'another page causes a problem. We may have to leave a 'forwarding address' on this page identifying the ne"v location of the record. And to ensure that space is ahvays available for this forwarding address, we would have to allocate some minimum space for each record, regardless of its length.

Large Records in Real Systems: In Sybase ASE, a record can be at most 1962 bytes. This limit is set by the 2KB log page size, since records are not allowed to be larger than a page. The exceptions to this rule are BLOBs and CLOBs, which consist of a set of bidirectionally linked pages. IBM DB2 and Microsoft SQL Server also do not allow records to span pages, although large objects are allowed to span pages and are handled separately from other data types. In DB2, record size is limited only by the page size; in SQL Server, a record can be at most 8KB, excluding LOBs. Informix and Oracle 8 allow records to span pages. Informix allows records to be at most 32KB, while Oracle has no maximum record size; large records are organized as a singly directed list.

- A record may grow so large that it no longer fits on *anyone* page. We have to deal with this condition by breaking a record into smaller records. The smaller records could be chained together-part of each smaller record is a pointer to the next record in the chain---to enable retrieval of the entire original record.

9.8 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Explain the term *memory hierarchy.* What are the differences between primary, secondary, and tertiary storage? Give examples of each. Which of these is *volatile*, and which are *persistent?* Why is persistent storage more important for a DBMS than, say, a program that generates prime numbers? (Section 9.1)
- Why are disks used so widely in a DBMS? What are their advantages over main memory and tapes? What are their relative disadvantages? (Section 9.1.1)
- What is a *disk block* or *page?* How are blocks arranged in a disk? How does this affect the time to access a block? Discuss *seek time*, *rotational delay*, and *transfer time.* (Section 9.1.1)
- Explain how careful placement of pages on the disk to exploit the geometry of a disk can minimize the seek time and rotational delay when pages are read sequentially. (Section 9.1.2)
- Explain what a RAID systenl is and how it improves performance and reliability. Discuss *striping* and its impact on performance and *redundancy* and its irnpact on reliability. What are the trade-offs between reliability

and performance in the different RAID organizations called *RAID levels'?* (Section 9.2)

- What is the role of the DBMS *disk space manager?* Why do database systems not rely on the operating system instead? (Section 9.3)
- Why does every page request in a DBMS go through the buffer manager? What is the *buffer poor?* What is the difference between a *frame* in a buffer pool, a *page* in a file, and a *block* on a disk? (Section 9.4)
- What information does the buffer manager maintain for each page in the buffer pool? ·What information is maintained for each frame? What is the significance of *pin_count* and the *dirty* flag for a page? Under what conditions can a page in the pool be *replaced?* Under what conditions must a replaced page be written back to disk? (Section 9.4)
- Why does the buffer manager have to replace pages in the buffer pool? How is a page chosen for replacement? What is *sequential flooding*, and what replacement policy causes it? (Section 9.4.1)
- A DBMS buffer manager can often predict the access pattern for disk pages. How does it utilize this ability to minimize I/O costs? Discuss *prefetching.* What is *forcing*, and why is it required to support the write-ahead log protocol in a DBMS? In light of these points, explain why database systems reimplement many services provided by operating systems. (Section 9.4.2)
- Why is the abstraction of a *file of records* important? How is the software in a DBMS layered to take advantage of this? (Section 9.5)
- What is a *heap file?* How are pages organized in a heap file? Discuss list versus directory organizations. (Section 9.5.1)
- Describe how records are arranged on a page. What is a *slot*, and how are slots used to identify records? How do slots ena.ble us to move records on a page withont altering the record's identifier? ·What arc the differences in page organizations for fixed-length and variable-length records? (Section 9.6)
- ·What are the differences in how fields are arranged within fixed-length and variable-length records? For variable-length records, explain how the array of offsets organization provides direct access to a specific field and supports *null* values. (Section 9.7)

EXERCISES

Exercise 9.1 What is the most important difference between a disk and a tape?

Exercise 9.2 Explain the terms *seek time, rotational delay,* and *transfer time.*

Exercise 9.3 Both disks and main memory support direct access to any desired location (page). On average, main memory accesses are faster, of course. What is the other important difference (from the perspective of the time required to access a desired page)?

Exercise 9.4 If you have a large file that is frequently scanned sequentially, explain how you would store the pages in the file on a disk.

Exercise 9.5 Consider a disk with a sector size of 512 bytes, 2000 tracks per surface, 50 sectors per track, five double-sided platters, and average seek time of 10 msec.

1. What is the capacity of a track in bytes? What is the capacity of each surface? What is the capacity of the disk?
2. How many cylinders does the disk have?
3. Give examples of valid block sizes. Is 256 bytes a valid block size? 2048? 51,200?
4. If the disk platters rotate at 5400 rpm (revolutions per minute), what is the maximum rotational delay?
5. If one track of data can be transferred per revolution, what is the transfer rate?

Exercise 9.6 Consider again the disk specifications from Exercise 9.5 and suppose that a block size of 1024 bytes is chosen. Suppose that a file containing 100,000 records of 100 bytes each is to be stored on such a disk and that no record is allowed to span two blocks.

1. How many records fit onto a block?
2. How many blocks are required to store the entire file? If the file is arranged sequentially on disk, how lllallY surfaces are needed?
3. How many records of 100 bytes each can be stored using this disk?
4. If pages are stored sequentially on disk, with page 1 on block 1 of track 1, what page is stored on block 1 of track 1 on the next disk surface? How would your answer change if the disk were capable of reading and writing from all heads in parallel?
5. What titne is required to read a file containing 100,000 records of 100 bytes each sequentially? Again, how \vould your answer change if the disk were capable of reading/writing from all heads in parallel (and the data was arranged optimally)?
6. What is the time required to read a file containing 100,000 records of 100 bytes each in a random order? To read a record, the block containing the recOl'd has to be fetched from disk. Assume that each block request incurs the average seek time and rotational delay.

Exercise 9.7 Explain what the buffer manager JIms! do to process a read request for a page. \Vhat happens if the requested page is in the pool but not pinned?

Exercise 9.8 When does a buffer manager write a page to disk?

Exercise 9.9 What does it mean to say that a page is *pinned* in the buffer pool? Who is responsible for pinning pages? \Vho is responsible for unpinning pages?

Exercise 9.10 'When a page in the bulTer pool is modified, how does the DBMS ensure that this change is propagated to disk? (Explain the role of the buffer manager as well as the modifier of the page.)

Exercise 9.11 \Vhat happens if a page is requested when all pages in the buffer pool are dirty?

Exercise 9.12 \Vhat is *sequential flooding* of the buffer pool?

Exercise 9.13 Name an important capability of a DBMS buffer manager that is not supported by a typical operating system's buffer manager.

Exercise 9.14 Explain the term *prefetching.* Why is it important?

Exercise 9.15 Modern disks often have their own main memory caches, typically about 1 MB, and use this to prefetch pages. The rationale for this technique is the empirical observation that, if a disk page is requested by some (not necessarily database!) application, 80% of the time the next page is requested as well. So the disk gambles by reading ahead.

1. Give a nontechnical reason that a DBMS may not want to rely on prefetching controlled by the disk.
2. Explain the impact on the disk's cache of several queries running concurrently, each scanning a different file.
3. Is this problem addressed by the DBMS buffer manager prefetching pages? Explain.
4. Modern disks support *segmented caches*, with about four to six segments, each of which is used to cache pages from a different file. Does this technique help, with respect to the preceding problem? Given this technique, does it matter whether the DBMS buffer manager also does prefetching?

Exercise 9.16 Describe two possible record formats. What are the trade-offs between them?

Exercise 9.17 Describe two possible page formats. What are the trade-offs between them?

Exercise 9.18 Consider the page format for variable-length records that uses a slot directory.

1. One approach to managing the slot directory is to use a maximum size (i.e., a maximum number of slots) and allocate the directory array when the page is created. Discuss the pros and cons of this approach with respect to the approach discussed in the text.
2. Suggest a modification to this page format that would allow us to sort records (according to the value in some field) without moving records and without changing the record ids.

Exercise 9.19 Consider the two internal organizations for heap files (using lists of pages and a directory of pages) discussed in the text.

1. Describe them briefly and explain the trade-offs. Which organization would you choose if records are variable in length?
2. Can you suggest a single page format to implement both internal file organizations'?

Exercise 9.20 Consider a list-based organizat.ion of the pages in a heap file in which two lists are maintained: a list of *all* pages in the file and a list of all pages with free space. In contrast, the list-based organizatioll discussed in the text maintains a list of full pages and a list of pages with free space.

1. What are the trade-offs, if any? Is one of them clearly superior?
2. For each of these organizations, describe a suitable page format.

Exercise **9.21** Modern disk drives store more sectors on the outer tracks than the inner tracks. Since the rotation speed is constant, the sequential data transfer rate is also higher on the outer tracks. The seek time and rotational delay are unchanged. Given this information, explain good strategies for placing files with the following kinds of access patterns:

1. Frequent, random accesses to a small file (e.g., catalog relations).
2. Sequential scans of a large file (e.g., selection from a relation with no index).
3. Random accesses to a large file via an index (e.g., selection from a relation via the index).
4. Sequential scans of a small file.

Exercise **9.22** Why do frames in the buffer pool have a pin count instead of a pin flag?

PROJECT-BASED EXERCISES

Exercise **9.23** Study the public interfaces for the disk space manager, the buffer manager, and the heap file layer in Minibase.

1. Are heap files with variable-length records supported?
2. What page format is used in Minibase heap files?
3. What happens if you insert a record whose length is greater than the page size?
4. How is free space handled in Minibase?

BIBLIOGRAPHIC NOTES

Salzberg [648] and Wiederhold [776] discuss secondary storage devices and file organizations in detail.

RAID was originally proposed by Patterson, Gibson, and Katz [587]. The article by Chen et al. provides an excellent survey of RAID [171] . Books about RAID include Gibson's dissertation [317] and the publications from the RAID Advisory Board [605].

The design and implementation of storage managers is discussed in [65, 133, 219, 477, 718]. With the exception of [219], these systems emphasize *extensibility*, anel the papers contain much of interest from that stanelpoint as well. Other papers that cover storage management issues in the context of significant implemented prototype systems are [480] and [588]. The Dali storage Inanager, which is optimized for main memory databases, is described in [406]. Three techniques for ilnplementing long fields are compared in [96]. The impact of processor cache misses 011 DBMS performallce has received attention lately, as complex queries have become increasingly CPU-intensive. [33] studies this issue, and shows that performance can be significantly improved by using a new arrangement of records within a page, in which records on a page are stored in a column-oriented format (all field values for the first attribute followed by values for the second attribute, etc.).

Stonebraker discusses operating systems issues in the context of databases in [715]. Several buffer management policies for database systems are compared in [181]. Buffer management is also studied in [119, 169, 2G1, 235].

10

TREE-STRUCTURED INDEXING

- What is the intuition behind tree-structured indexes? Why are they good for range selections?
- How does an ISAM index handle search, insert, and delete?
- How does a B+ tree index handle search, insert, and delete?
- What is the impact of duplicate key values on index implementation'?
- What is key compression, and why is it important?
- What is bulk-loading, and why is it important?
- What happens to record identifiers when dynamic indexes are updated? How does this affect clustered indexes?

Key concepts: ISAM, static indexes, overflow pages, locking issues; B+ trees, dynamic indexes, balance, sequence sets, node format; B+ tree insert operation, node splits, delete operation, merge versus redistribution, minimum occupancy; duplicates, overflow pages, including rids in search keys; key compression; bulk-loading; effects of splits on rids in clustered indexes.

One that would have the fruit must climb the tree.

Thomas Fuller

We now consider two index data structures, called ISAM and B+ trees, based on tree organizations. These structures provide efficient support for range searches, including sorted file scans as a special case. Unlike sorted files, these

index structures support efficient insertion and deletion. They also provide support for equality selections, although they are not as efficient in this case as hash-based indexes, which are discussed in Chapter 11.

An ISAJVI[1] tree is a static index structure that is effective when the file is not frequently updated, but it is unsuitable for files that grow and shrink a lot. \Ve discuss ISAM in Section 10.2. The B+ tree is a dynamic structure that adjusts to changes in the file gracefully. It is the most widely used index structure because it adjusts well to changes and supports both equality and range queries. We introduce B+ trees in Section 10.3. We cover B+ trees in detail in the remaining sections. Section 10.3.1 describes the format of a tree node. Section *IOA* considers how to search for records by using a B+ tree index. Section 10.5 presents the algorithm for inserting records into a B+ tree, and Section 10.6 presents the deletion algorithm. Section 10.7 discusses how duplicates are handled. We conclude with a discussion of some practical issues concerning B+ trees in Section 10.8.

Notation: In the ISAM and B+ tree structures, leaf pages contain *data entries*, according to the terminology introduced in Chapter 8. For convenience, we denote a data entry with search key value k as $k*$. Non-leaf pages conta.in *index entries* of the form *(search key value, page id)* and are used to direct the sea.rch for a desired data entry (which is stored in some leaf). We often simply use *entr'Y* where the context makes the nature of the entry (index or data) clear.

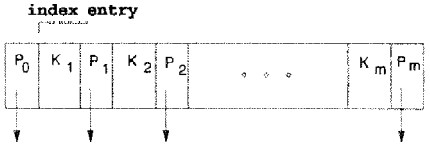
10.1 INTUITION FOR TREE INDEXES

Consider a file of Students recorcls sorted by *gpa*. To answer a range selection such as "Find all students with a gpa higher than 3.0," we must identify the first such student by doing a binary search of the file and then scan the file from that point on. If the file is large, the initial binary search can be quite expensive, since cost is proportional to the number of pages fetched; can we improve upon this method?

OIle idea is to create a second file with OIle record per page in the original (data) file, of the form *(first key on page, pointer to page)*, again sortecl by the key attribute (which is *gpa* in our example). The format of a page in the second *index* file is illustrated in Figure 10.1.

We refer to pairs of the form *(key, pointer)* as *index entries* or just *entries* when the context is clear. Note that each index page contains OIle pointer more than

[1] ISAM stands for Indexed Sequential Access Method.



FigUl'e 10.1 Format of an Index Page

the number of keys—each key serves as a *separator* for the contents of the pages pointed to by the pointers to its left and right.

The simple index file data structure is illustrated in Figure 10.2.

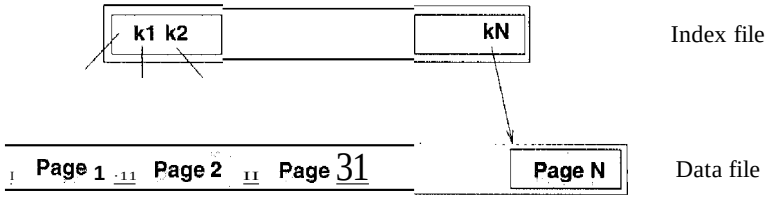


Figure 10.2 One-Level Index Structure

We can do a binary search of the index file to identify the page containing the first key *(gpo.)* value that satisfies the range selection (in our example, the first student with *gpa* over 3.0) and follow the pointer to the page containing the first data. record with that key value. We can then scan the data file sequentially from that point on to retrieve other qualifying records. This example uses the index to find the first data page containing a Students record with *gpa* greater than 3.0, and the data file is scanned from that point on to retrieve other such Students records.

Because the size of an entry in the index file (key value and page icl) is likely to be much smaller than the size of a page, and only one such entry exists per page of the data file, the index file is likely to be much smaller than the data file; therefore, a binary search of the index file is much faster than a binary search of the data file. However, a binary search of the index file could still be fairly expensive, and the index file is typically still large enough to make inserts and deletes expensive.

The potential large size of the index file motivates the tree indexing idea: Why not apply the previous step of building an auxiliary structure all the collection of *index* records and so on recursively until the smallest auxiliary structure fits on one page? This repeated construction of a one-level index leads to a tree structure with several levels of non-leaf pages.

As we observed in Section 8.3.2, the power of the approach comes from the fact that locating a record (given a search key value) involves a traversal from the root to a leaf, with one I/O (at most; some pages, e.g.) the root, are likely to be in the buffer pool) per level. Given the typical fan-out value (over 100), trees rarely have more than 3–4 levels.

The next issue to consider is how the tree structure can handle inserts and deletes of data entries. Two distinct approaches have been used, leading to the ISAM and B+ tree data structures, which we discuss in subsequent sections.

10.2 INDEXED SEQUENTIAL ACCESS METHOD (ISAM)

The ISAM data structure is illustrated in Figure 10.3. The data entries of the ISAM index are in the leaf pages of the tree and additional *overflow* pages chained to some leaf page. Database systems carefully organize the layout of pages so that page boundaries correspond closely to the physical characteristics of the underlying storage device. The ISAM structure is completely static (except for the overflow pages, of which it is hoped, there will be few) and facilitates such low-level optimizations.

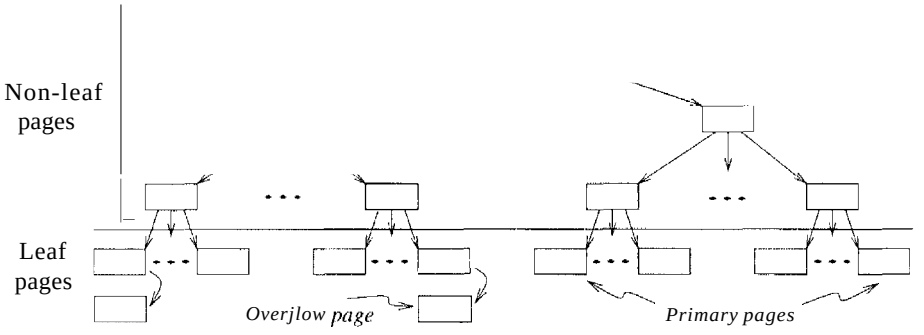


Figure 10.3 ISAM Index Structure

Each tree node is a disk page, and all the data resides in the leaf pages. This corresponds to an index that uses Alternative (1) for data entries, in terms of the alternatives described in Chapter 8; we can create an index with Alternative (2) by storing the data records in a separate file and storing $\langle key, rid \rangle$ pairs in the leaf pages of the ISAM index. When the file is created, all leaf pages are allocated sequentially and sorted on the search key value. (If Alternative (2) or (3) is used, the data records are created and sorted before allocating the leaf pages of the ISAM index.) The non-leaf level pages are then allocated. If there are several inserts to the file subsequently, so that more entries are inserted into a leaf than will fit onto a single page, additional pages are needed because the

index structure is static. These additional pages are allocated from an overflow area. The allocation of pages is illustrated in Figure 10.4.

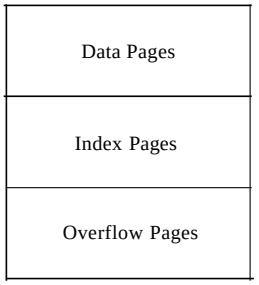


Figure 10.4 Page Allocation in ISAM

The basic operations of insertion, deletion, and search are all quite straightforward. For an equality selection search, we start at the root node and determine which subtree to search by comparing the value in the search field of the given record with the key values in the node. (The search algorithm is identical to that for a B+ tree; we present this algorithm in more detail later.) For a range query, the starting point in the data (or leaf) level is determined similarly, and data pages are then retrieved sequentially. For inserts and deletes, the appropriate page is determined as for a search, and the record is inserted or deleted with overflow pages added if necessary.

The following example illustrates the ISAM index structure. Consider the tree shown in Figure 10.5. All searches begin at the root. For example, to locate a record with the key value 27, we start at the root and follow the left pointer, since 27 < 40. We then follow the middle pointer, since 20 <= 27 < 33. For a range sea,rch, we find the first qualifying data entry as for an equality selection and then retrieve primary leaf pages sequentially (also retrieving overflow pages as needed by following pointers from the primary pages). The primary leaf pages are assumed to be allocated sequentially this assumption is reasonable because the number of such pages is known when the tree is created and does not change subsequently under inserts and deletes-and so no 'next leaf page' pointers are needed.

We assume that each leaf page can contain two entries. If we now insert a record with key value 23, the entry 23* belongs in the second data page, which already contains 20* and 27* and has no more space. We deal with this situation by adding an *overflow* page and putting 23* in.the overflow page. Chains of overflow pages can easily develop. For instance, inserting 48*, 41*, and 42* leads to an overflow chain of two pages. The tree of Figure 10.5 with all these insertions is shown ill Figure 10.6.

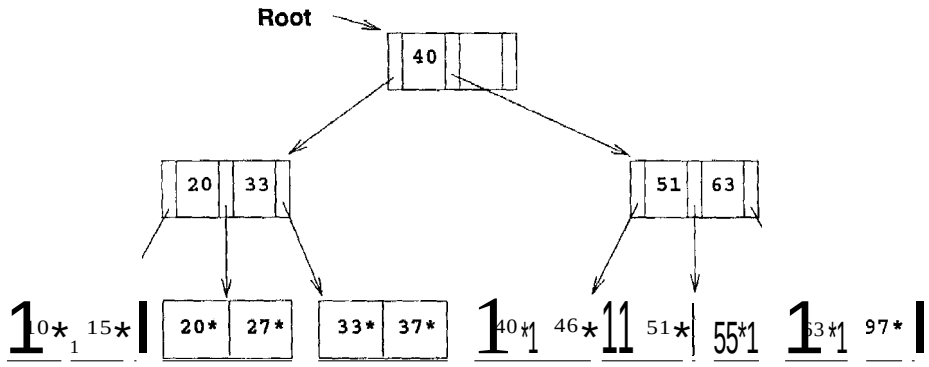


Figure 10.5 Sample ISAM Tree

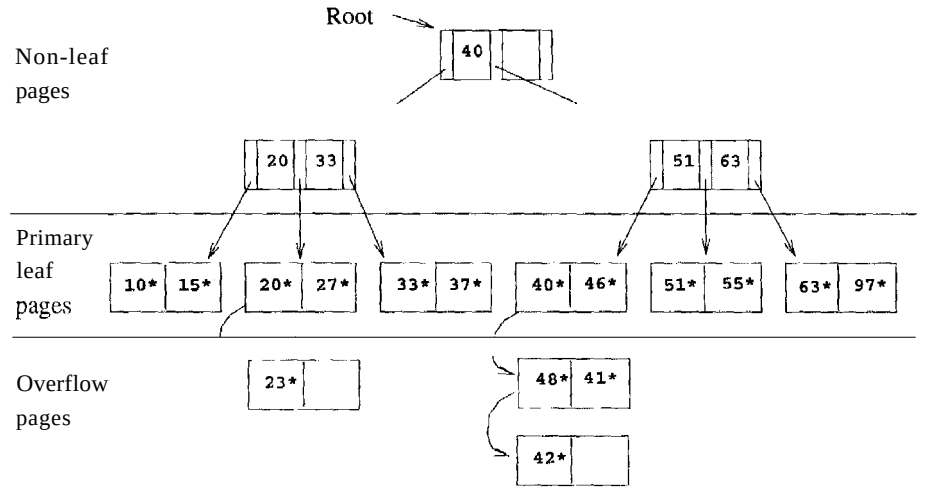


Figure 10.6 ISAM Tree after Inserts

The deletion of an entry $k*$ is handled by simply removing the entry. If this entry is on an overflow page and the overflow page becomes empty, the page can be removed. If the entry is on a primary page and deletion makes the primary page empty, the simplest approach is to simply leave the empty primary page as it is; it serves as a placeholder for future insertions (and possibly lloll-empty overflow pages, because we do not move records from the overflow pages to the primary page when deletions on the primary page create space). Thus, the number of primary leaf pages is fixed at file creation time.

10.2.1 Overflow Pages, Locking Considerations

Note that, once the ISAM file is created, inserts and deletes affect only the contents of leaf pages. A consequence of this design is that long overflow chains could develop if a number of inserts are made to the same leaf. These chains can significantly affect the time to retrieve a record because the overflow chain has to be searched as well when the search gets to this leaf. (Although data in the overflow chain can be kept sorted, it usually is not, to make inserts fast.) To alleviate this problem, the tree is initially created so that about 20 percent of each page is free. However, once the free space is filled in with inserted records, unless space is freed again through deletes, overflow chains can be eliminated only by a complete reorganization of the file.

The fact that only leaf pages are modified also has an important advantage with respect to concurrent access. When a page is accessed, it is typically 'locked' by the requestor to ensure that it is not concurrently modified by other users of the page. To modify a page, it must be locked in 'exclusive' mode, which is permitted only when no one else holds a lock on the page. Locking can lead to queues of users (*transactions*, to be more precise) waiting to get access to a page. Queues can be a significant performance bottleneck, especially for heavily accessed pages near the root of an index structure. In the ISAM structure, since we know that index-level pages are never modified, we can safely omit the locking step. Not locking index-level pages is an important advantage of ISAM over a dynamic structure like a B+ tree. If the data distribution and size are relatively static, which means overflow chains are rare, ISAM might be preferable to B+ trees due to this advantage.

10.3 B+ TREES: A DYNAMIC INDEX STRUCTURE

A static structure such as the ISAM index suffers from the problem that long overflow chains can develop as the file grows, leading to poor performance. This problem motivated the development of more flexible, dynamic structures that adjust gracefully to inserts and deletes. The B+ tree search structure, which is widely llsed, is a balanced tree in which the internal nodes direct the search

and the leaf nodes contain the data entries. Since the tree structure grows and shrinks dynamically, it is not feasible to allocate the leaf pages sequentially as in ISAM, where the set of primary leaf pages was static. To retrieve all leaf pages efficiently, we have to link them using page pointers. By organizing them into a doubly linked list, we can easily traverse the sequence of leaf pages (sometimes called the sequence set) in either direction. This structure is illustrated in Figure 10.7.[2]

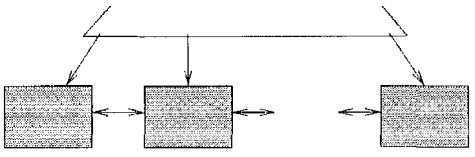


Figure 10.7 Structure of a B+ Tree

The following are some of the main characteristics of a B+ tree:

- Operations (insert, delete) on the tree keep it balanced.
- A minimum occupancy of 50 percent is guaranteed for each node except the root if the deletion algorithm discussed in Section 10.6 is implemented. However, deletion is often implemented by simply locating the data entry and removing it, without adjusting the tree as needed to guarantee the 50 percent occupancy, because files typically grow rather than shrink.
- Searching for a record requires just a traversal from the root to the appropriate leaf. We refer to the length of a path from the root to a leaf—any leaf, because the tree is balanced—as the height of the tree. For example, a tree with only a leaf level and a single index level, such as the tree shown in Figure 10.9, has height 1, and a tree that has only the root node has height 0. Because of high fan-out, the height of a B+ tree is rarely more than 3 or 4.

We will study B+ trees in which every node contains m entries, where $d \leq m \leq 2d$. The value d is a parameter of the B+ tree, called the order of the

[2] If the tree is created by *bulk_looding* (see Section 10.8.2) an existing data set, the sequence set can be made physically sequential, but this physical ordering is gradually destroyed as new data is added and deleted over time.

tree, and is a measure of the capacity of a tree node. The root node is the only exception to this requirement on the number of entries; for the root, it is simply required that $1 \leq m \leq 2d$.

If a file of records is updated frequently and sorted access is important, maintaining a B+ tree index with data records stored as data entries is almost always superior to maintaining a sorted file. For the space overhead of storing the index entries, we obtain all the advantages of a sorted file plus efficient insertion and deletion algorithms. B+ trees typically maintain 67 percent space occupancy. B+ trees are usually also preferable to ISAM indexing because inserts are handled gracefully without overflow chains. However, if the dataset size and distribution remain fairly static, overflow chains may not be a major problem. In this case, two factors favor ISAM: the leaf pages are allocated in sequence (making scans over a large range more efficient than in a B+ tree, in which pages are likely to get out of sequence on disk over time, even if they were in sequence after bulk-loading), and the locking overhead ofISAM is lower than that for B+ trees. As a general rule, however, B+ trees are likely to perform better than ISAM.

10.3.1 Format of a Node

The format of a node is the same as for ISAM and is shown in Figure 10.1. Non-leaf nodes with m *index entries* contain $m+1$ pointers to children. Pointer *Pi* points to a subtree in which all key va.lues K are such that $K_i \leq K < \mathrm{K_i{+}_1}$. As special cases, *Po* points to a tree in which all key values are less than Kl' and P_m points to a tree in which all key values are greater than or equal to K_m. For leaf nodes, entries arc denoted as *k**, as usual. Just as in ISAM, leaf nodes (and *only* leaf nodes!) contain *data entries.* In the common case that Alternative (2) or (3) is used, leaf entries are *(K,I(K)* $\rangle$ pairs, just like non-leaf entries. Regardless of the alternative chosen for leaf entries, the leaf pages are chained together in a doubly linked list. Thus, the leaves form a sequence, which can be used to answer range queries efficiently.

The reader should carefully consider how such a node organization can be achieved using the record formats presented in Section 9.7; after all, each key pointer pair can be thought of as a record. If the field being indexed is of fixed length, these index entries will be of fixed length; otherwise, we have variable-length records. In either case the B+ tree can itself be viewed as a file of records. If the leaf pages do not contain the actual data records, then the 13+ tree is indeed a file of records that is distinct from the file that contains the data. If the leaf pages contain data records, then a file contains the 13+ tree as well as the data.

10.4 SEARCH

The algorithm for search finds the leaf node in which a given data entry belongs. A pseudocode sketch of the algorithm is given in Figure 10.8. We use the notation **ptr* to denote the value pointed to by a pointer variable *ptr* and & *(value)* to denote the address of *value*. Note that finding *i* in *tTcc_seaTch* requires us to search within the node, which can be done with either a linear search or a binary search (e.g., depending on the number of entries in the node).

In discussing the search, insertion, and deletion algorithms for B+ trees, we assume that there are no *duplicates*. That is, no two data entries are allowed to have the same key value. Of course, duplicates arise whenever the search key does not contain a candidate key and must be dealt with in practice. We consider how duplicates can be handled in Section 10.7.

```
fune find (search key value K) returns nodepointer
// Given a search key value, finds its leaf node
return tree_search(root, K);                    // searches from root
endfune

fune tree_search (nodepointer, search key value K) returns nodepointer
// Searches tree for entry
if *nodepointer is a leaf, return nodepointer;
else,
     if K < K_1 then return tree_search(P_0, K);
     else,
          if K ≥ K_m then return tree_search(P_m, K);    // 171 = # entries
          else,
               find i such that K_i ≤ K < K_{i+1};
               return tree_search(P_i, K)
endfune
```

Figure 10.8 Algorithm for Search

Consider the sample B+ tree shown in Figure 10.9. This B+ tree is of order d=2. That is, each node contains between 2 and 4 entries. Each non--leaf entry is a $\langle key\ value, nodepointer \rangle$ pair; at the leaf level, the entries are data records that we denote by $k*$. To search for entry 5*, we follow the left-most child pointer, since 5 < 13. To search for the entries 14* or 15*, we follow the second pointer, since $13 \leq 14 < 17$, and $13 \leq 15 < 17$. (We do not find 15* on the appropriate leaf and can conclude that it is not present in the tree.) To find 24*, we follow the fourth child pointer, since $24 \leq 24 < 30$.

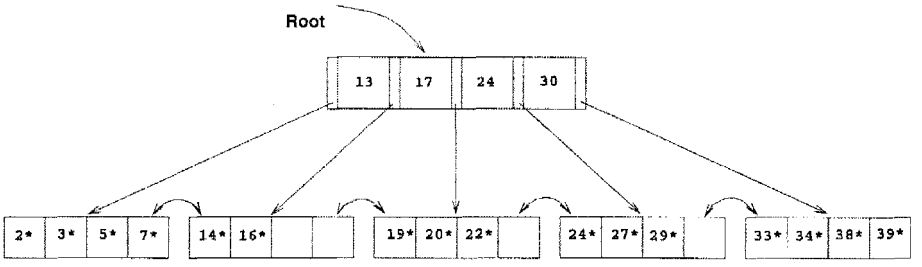


Figure 10.9 Example of a B+ Tree, Order d=2

10.5 INSERT

The algorithm for insertion takes an entry, finds the leaf node where it belongs, and inserts it there. Pseudocode for the B+ tree insertion algorithm is given in Figure HUG. The basic idea behind the algorithm is that we recursively insert the entry by calling the insert algorithm on the appropriate child node. Usually, this procedure results in going down to the leaf node where the entry belongs, placing the entry there, and returning all the way back to the root node. Occasionally a node is full and it must be split. When the node is split, an entry pointing to the node created by the split must be inserted into its parent; this entry is pointed to by the pointer variable *newchildentry*. If the (old) root is split, a new root node is created and the height of the tree increases by 1.

To illustrate insertion, let us continue with the sample tree shown in Figure 10.9. If we insert entry 8*, it belongs in the left-most leaf, which is already full. This insertion causes a split of the leaf page; the split pages are shown in Figure 10.11. The tree must now be adjusted to take the new leaf page into account, so we insert an entry consisting of the pair (5, *pointer to new page)* into the parent node. Note how the key 5, which discriminates between the split leaf page and its newly created sibling, is 'copied up.' We cannot just 'push up' 5, because every data entry must appear in a leaf page.

Since the parent node is also full, another split occurs. In general we have to split a non-leaf node when it is full, containing $2d$ keys and $2d+1$ pointers, and we have to add another index entry to account for a child split. We now have $2d+1$ keys and $2d+2$ pointers, yielding two minimally full non-leaf nodes, each containing d keys and $d+1$ pointers, and an extra key, which we choose to be the 'middle' key. This key and a pointer to the second non-leaf node constitute an index entry that must be inserted into the parent of the split non-leaf node. The middle key is thus 'pushed up' the tree, in contrast to the case for a split of a leaf page.

```
proc inseTt (nodepointel', entry, newchildentry)
// InseTts entry into subtree with TOot '*nodepointer'; degree is d;
//'newchildentry' null initially, and null on return unless child is split

if *nodepointer is a non-leaf node, say N,
    find'i such that K_i ≤ entry's key value < J(i+1;    // choose subtree
    insert(P_i, entry, newchildentry);          // recurs'ively, insert entry
    if newchildentry is null, return;          // usual case; didn't split child
    else,                  // we split child, must insert *newchildentry in N
        if N has space,                                        // usual case
            put *newchildentry on it, set newchildentry to null, return;
        else,                  // note difference wrt splitting of leaf page!
            split N:          // 2d + 1 key values and 2d + 2 nodepointers
            first d key values and d + 1 nodepointers stay,
            last d keys and d + 1 pointers move to new node, N2;
            // *newchildentry set to guide searches between N and N2
            newchildentry = & ((smallest key value on N2,
                                pointer to N2));
            if N is the root,                  // root node was just split
                create new node with (pointer to N, *newchildentry);
                make the tree's root-node pointer point to the new node;
            return;

if *nodepointer is a leaf node, say L,
    if L has space,                                            // usual case
    put entry on it, set newchildentry to null, and return;
    else,                                  // once in a while, the leaf is full
        split L: first d entries stay, rest move to brand new node L2;
        newchildentry = & ((smallest key value on L2, pointer to L2));
        set sibling pointers in L and L2;
        return;
endproc
```

Figure 10.1.0 Algorithrn for Insertion into B+ Tree of Order *d*

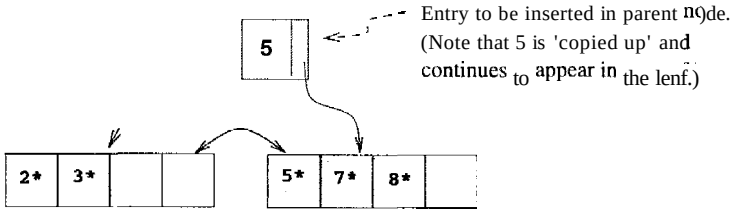


Figure 10.11 Split Leaf Pages during Insert of Entry 8*

The split pages in our example are shown in Figure 10.12. The index entry pointing to the new non-leaf node is the pair ⟨*17, pointer to new index-level page*⟩; note that the key value 17 is 'pushed up' the tree, in contrast to the splitting key value 5 in the leaf split, which was 'copied up.'

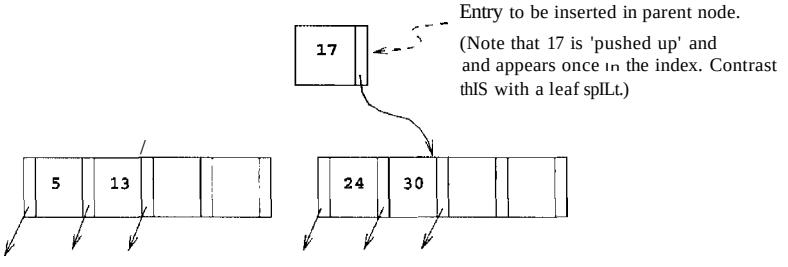


Figure 10.12 Split Index Pages during Insert of Entry 8*

The difference in handling leaf-level and index-level splits arises from the B+ tree requirement that all data entries $k*$ must reside in the leaves. This requirement prevents us from 'pushing up' 5 and leads to the slight redundancy of having some key values appearing in the leaf level as well as in some index level. However, range queries can be efficiently answered by just retrieving the sequence of leaf pages; the redundancy is a small price to pay for efficiency. In dealing with the index levels, we have more flexibility, and we 'push up' 17 to avoid having two copies of 17 in the index levels.

Now, since the split node was the old root, we need to create a new root node to hold the entry that distinguishes the two split index pages. The tree after completing the insertion of the entry 8* is shown in Figure 10.13.

One variation of the insert algorithm tries to redistribute entries of a node N with a sibling before splitting the node; this improves average occupancy. The sibling of a node N, in this context, is a node that is immediately to the left or right of N *and has the same parent as N.*

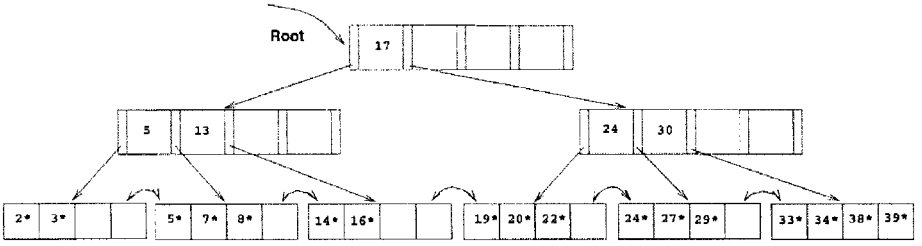


Figure 10.13 B+ Tree after Inserting Entry 8*

To illustrate redistribution, reconsider insertion of entry 8* into the tree shown in Figure 10.9. The entry belongs in the left-most leaf, which is full. However, the (only) sibling of this leaf node contains only two entries and can thus accommodate more entries. We can therefore handle the insertion of 8* with a redistribution. Note how the entry in the parent node that points to the second leaf has a new key value; we 'copy up' the new low key value on the second leaf. This process is illustrated in Figure 10.14.

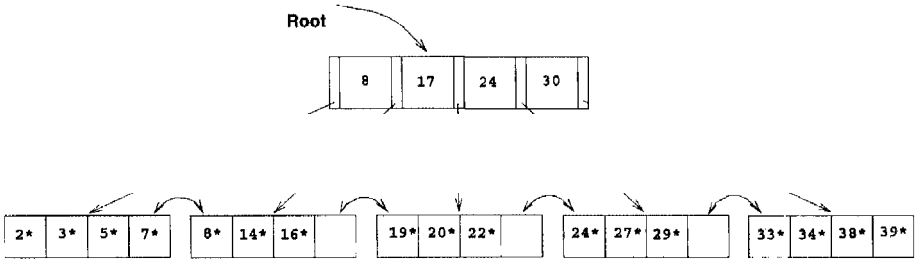


Figure 10.14 B+ Tree after Inserting Entry 8* Using Redistribution

To determine whether redistribution is possible, we have to retrieve the sibling. If the sibling happens to be full, we have to split the node anyway. On average, checking whether redistribution is possible increases I/O for index node splits, especially if we check both siblings. (Checking whether redistribution is possible may reduce I/O if the redistribution succeeds whereas a split propagates up the tree, but this case is very infrequent.) If the file is growing, average occupancy will probably not be affected much even if we do not redistribute. Taking these considerations into account, *not* redistributing entries at non-leaf levels usually pays off.

If a split occurs at the leaf level, however, we have to retrieve a neighbor to adjust the previous and next-neighbor pointers with respect to the newly created leaf node. Therefore, a limited form of redistribution makes sense: If a leaf node is full, fetch a neighbor node; if it has space and has the same parent,

redistribute the entries. Othenvise (the neighbor has different parent, Le., it is not a sibling, or it is also full) split the leaf node and a,djust the previous and next-neighbor pointers in the split node, the newly created neighbor, and the old neighbor.

10.6 DELETE

The algorithm for deletion takes an entry, finds the leaf node where it belongs, and deletes it. Pseudocode for the B+ tree deletion algorithm is given in Figure 10.15. The basic idea behind the algorithm is that we recursively delete the entry by calling the delete algorithm on the appropriate child node. We usually go down to the leaf node where the entry belongs, remove the entry from there, and return all the way back to the root node. Occasionally a node is at minimum occupancy before the deletion, and the deletion causes it to go below the occupancy threshold. When this happens, we must either redistribute entries from an adjacent sibling or merge the node with a sibling to maintain minimum occupancy. If entries are redistributed between two nodes, their parent node must be updated to reflect this; the key value in the index entry pointing to the second node must be changed to be the lowest search key in the second node. If two nodes are merged, their parent must be updated to reflect this by deleting the index entry for the second node; this index entry is pointed to by the pointer variable *oldchildentry* when the delete call returns to the parent node. If the last entry in the root node is deleted in this manner because one of its children was deleted, the height of the tree decreases by 1.

To illustrate deletion, let us consider the sample tree shown in Figure 10.13. To delete entry 19*, we simply remove it from the leaf page on which it appears, and we are done because the leaf still contains two entries. If we subsequently delete 20*, however, the leaf contains only one entry after the deletion. The (only) sibling of the leaf node that contained 20* has three entries, and we can therefore deal with the situation by redistribution; we move entry 24* to the leaf page that contained 20* and copy up the new splitting key (27, which is the new low key value of the leaf from which we borrowed 24*) into the parent. This process is illustrated in Figure 10.16.

Suppose that we now delete entry 24*. The affected leaf contains only one entry (22*) after the deletion, and the (only) sibling contains just two entries (27* and 29*). Therefore, we cannot redistribute entries. However, these two leaf nodes together contain only three entries and can be merged. \Vhile merging, we can 'toss' the entry (*27, pointer' to second leaf page)*) in the parent, which pointed to the second leaf page, because the second leaf page is elnpty after the merge and can be discarded. The right subtree of Figure 10.16 after this step in the deletion of entry 24* is shown in Figure 10.17.

```
proc delete (parentpointer, nodepointer, entry, oldchiIdentry)
// Deletes entry from s'ubtree with TOot '*nodepointer'; degree is d;
// 'oldchildentry' null initially, and null upon return unless child deleted
if *nodepointer is a non-leaf node, say N,
    find i such that K_i ≤ entry's key value < K_i+I;    // choose subtree
    delete(nodepointer, Pi, entry, oldchildentry);       // recursive delete
    if oldchildentry is null, return;          // usual case: child not deleted
    else,                                      // we discarded child node (see discussion)
        remove *oldchildentry from N,          // next, check for underflow
        if N has entries to spare,                                  // usual case
            set oldchildentry to null, return; // delete doesn't go further
        else,                                  // note difference wrt merging of leaf pages!
            get a sibling S of N:              // parentpointer arg used to find S
            if S has extra entries,
                redistribute evenly between Nand S through parent;
                set oldchildentry to null, return;
            else, merge Nand S                 // call node on rhs M
                oldchildentry = & (current entry in parent for M);
                pull splitting key from parent down into node on left;
                move all entries from M to node on left;
                discard empty node M, return;

if *nodepointer is a leaf node, say L,
    if L has entries to spare,                                  // usual case
        remove entry, set oldchildentry to null, and return;
    else,                                  // once in a while, the leaf becomes underfull
        get a sibling S of L;              // parentpointer used to find S
        if S has extra entries,
            redistribute evenly between Land S;
            find entry in parent for node on right;          // call it M
            replace key value in parent entry by new low-key value in M;
            set oldchildentry to null, return;
        else, merge Land S                 // call node on rhs M
            oldchildentry = & (current entry in parent for M);
            move all entries from M to node on left;
            discard empty node M, adjust sibling pointers, return;
endproc
```

Figure 10.15 Algorithm for Deletion from B+ Tree of Order d

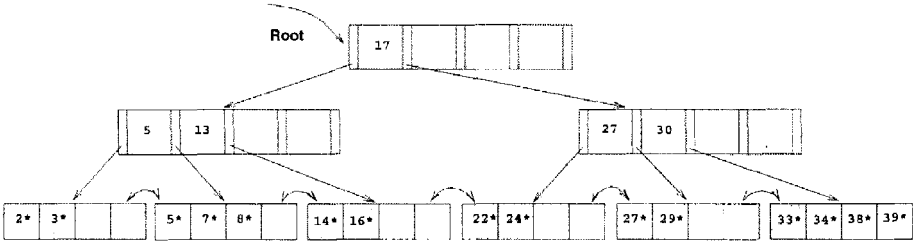


Figure 10.16 B+ Tree after Deleting Entries 19* and 20*

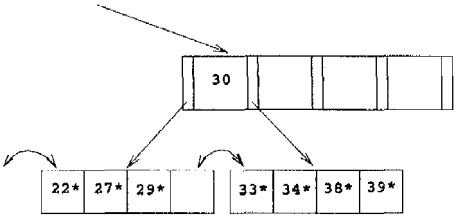


Figure 10.17 Partial B+ Tree during Deletion of Entry 24*

Deleting the entry (27, *pointer to second leaf page)* has created a non-Ieaf-Ievel page with just one entry, which is below the minimum of $d = 2$. To fix this problem, we must either redistribute or merge. In either case, we must fetch a sibling. The only sibling of this node contains just two entries (with key values 5 and 13), and so redistribution is not possible; we must therefore merge.

The situation when we have to merge two non-leaf nodes is exactly the opposite of the situation when we have to split a non-leaf node. We have to split a non-leaf node when it contains $2d$ keys and $2d + 1$ pointers, and we have to add another key--pointer pair. Since we resort to merging two non-leaf nodes only when we cannot redistribute entries between them, the two nodes must be minimally full; that is, each must contain d keys and $d + 1$ pointers prior to the deletion. After merging the two nodes and removing the key--pointer pair to be deleted, we have $2d - 1$ keys and $2d + 1$ pointers: Intuitively, the left-most pointer on the second merged node lacks a key value. To see what key value must be combined with this pointer to create a complete index entry, consider the parent of the two nodes being merged. The index entry pointing to one of the merged nodes must be deleted from the parent because the node is about to be discarded. The key value in this index entry is precisely the key value we need to complete the new merged node: The entries in the first node being merged, followed by the splitting key value that is 'pulled down' from the parent, followed by the entries in the second non-leaf node gives us a total of $2d$ keys and $2d + 1$ pointers, which is a full non-leaf node. Note how the splitting

key value in the parent is pulled down, in contrast to the case of merging two leaf nodes.

Consider the merging of two non-leaf nodes in our example. Together, the non-leaf node and the sibling to be merged contain only three entries, and they have a total of five pointers to leaf nodes. To merge the two nodes, we also need to pull down the index entry in their parent that currently discriminates between these nodes. This index entry has key value 17, and so we create a new entry *(17, left-most child pointer in sibling).* Now we have a total of four entries and five child pointers, which can fit on one page in a tree of order $d = 2$. Note that pulling down the splitting key 17 means that it will no longer appear in the parent node following the merge. After we merge the affected non-leaf node and its sibling by putting all the entries on one page and discarding the empty sibling page, the new node is the only child of the old root, which can therefore be discarded. The tree after completing all these steps in the deletion of entry 24* is shown in Figure 10.18.

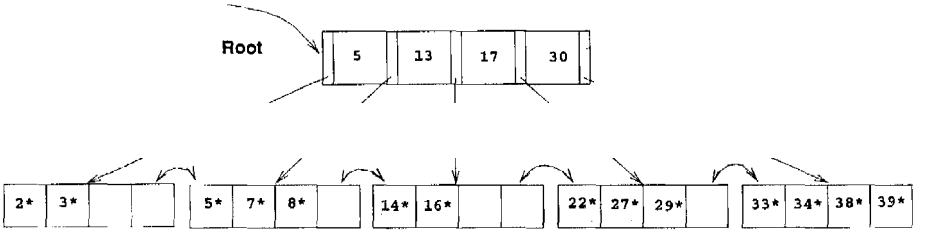


Figure 10.18 B+ Tree after Deleting Entry 24*

The previous examples illustrated redistribution of entries across leaves and merging of both leaf-level and non-leaf-level pages. The remaining case is that of redistribution of entries between non-leaf-level pages. To understand this case, consider the intermediate right subtree shown in Figure 10.17. We would arrive at the same intermediate right subtree if we try to delete 24* from a tree similar to the one shown in Figure 10.16 but with the left subtree and root key value as shown in Figure 10.19. The tree in Figure 10.19 illustrates an intermediate stage during the deletion of 24*. (Try to construct the initial tree.)

In contrast to the case when we deleted 24* from the tree of Figure 10.16, the non-leaf level node containing key value 30 now has a sibling that can spare entries (the entries with key values 17 and 20). We move these entries[3] over from the sibling. Note that, in doing so, we essentially push them through the

[3]It is sufficient to move over just the entry with key value 20, hut we are moving over two entries so illustrate what happens when several entries are redistributed.

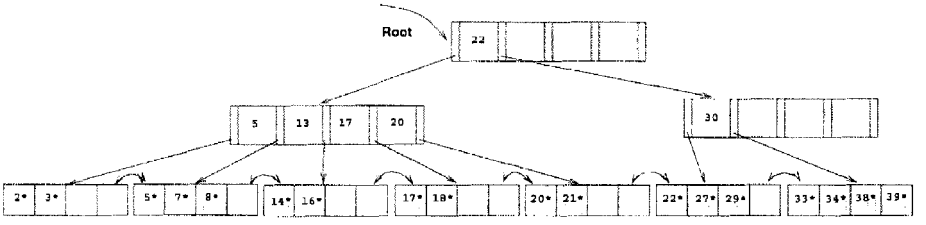


Figure 10.19 A B+ Tree during a Deletion

splitting entry in their parent node (the root), which takes care of the fact that 17 becomes the new low key value on the right and therefore must replace the old splitting key in the root (the key value 22). The tree with all these changes is shown in Figure 10.20.

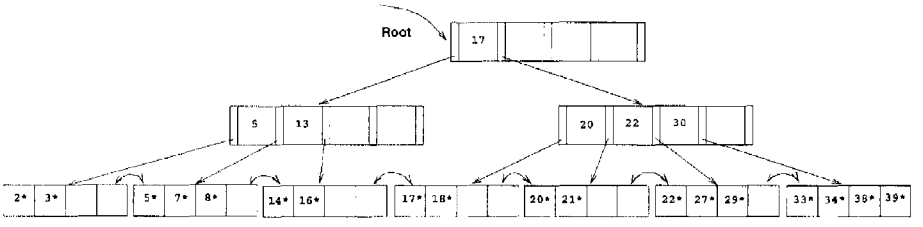


Figure 10.20 B+ Tree after Deletion

In concluding our discussion of deletion, we note that we retrieve only one sibling of a node. If this node has spare entries, we use redistribution; otherwise, we merge. If the node has a second sibling, it may be worth retrieving that sibling as well to check for the possibility of redistribution. Chances are high that redistribution is possible, and unlike merging, redistribution is guaranteed to propagate no further than the parent node. Also, the pages have more space on them, which reduces the likelihood of a split on subsequent insertions. (Remember, files typically grow, not shrink!) However, the number of times that this case arises (the node becomes less than half-full and the first sibling cannot spare an entry) is not very high, so it is not essential to implement this refinement of the basic algorithm that we presented.

10.7 DUPLICATES

The search, insertion, and deletion algorithms that we presented ignore the issue of duplicate keys, that is, several data entries with the same key value. We now discuss how duplica.tes can be handled.

Duplicate Handling in **Commercial** Systems: In a clustered index in Sybase ASE, the data rows are maintained in sorted order onthe page and in the eollection of data pages. The data pages are bidireetionally linked in sort order. Rows with duplicate keys are inserted into (or deleted from) the ordered set of rows. This may result in overflow pages of rows with duplieate keys being inserted into the page chain or empty overflow pages removed from the page chain. Insertion or deletion of a duplicate key does not affect the higher index levels unless a split or merge ofa non.-overflow page occurs. In IBM DB2, Oracle 8, and Microsoft SQL Server, duplicates are handled by adding a row id if necessary to eliminate duplicate key values.

The basic search algorithm assumes that all entries with a given key value reside on a single leaf page. One way to satisfy this assumption is to use *overflow pages* to deal with duplicates. (In ISAM, of course, we have overflow pages in any case, and duplicates are easily handled.)

Typically, however, we use an alternative approach for duplicates. We handle them just like any other entries and several leaf pages may contain entries with a given key value. To retrieve all data entries with a given key value, we must search for the *left-most* data entry with the given key value and then possibly retrieve more than one leaf page (using the leaf sequence pointers). Modifying the search algorithm to find the left-most data entry in an index with duplicates is an interesting exercise (in fact, it is Exercise 10.11).

One problem with this approach is that, when a record is deleted, if we use Alternative (2) for data entries, finding the corresponding data entry to delete in the B+ tree index could be inefficient because we may have to check several duplicate entries *(key, rid)* with the same *key* value. This problem can be addressed by considering the *rid* value in the data entry to be *part of the search key,* for purposes of positioning the data entry in the tree. This solution effectively turns the index into a *unique* index (i.e" no duplicates), Remember that a search key can be any sequence of fields in this variant, the rid of the data record is essentially treated as another field while constructing the search key.

Alternative (3) for data entries leads to a natural solution for duplicates, but if we have a large number of duplicates, a single data entry could span multiple pages. And of course, when a data record is deleted, finding the rid to delete from the corresponding data entry can be inefficient, The solution to this problem is similar to the one discussed previously for Alternative (2): We can

maintain the list of rids within each data entry in sorted order (say, by page number and then slot number if a rid consists of a page id and a slot id).

10.8 B+ TREES IN PRACTICE

In this section we discuss several important pragmatic issues.

10.8.1 Key Compression

The height of a B+ tree depends on the *number of data entries* and the *size of index entries.* The size of index entries determines the number of index entries that will fit on a page and, therefore, the *fan-out* of the tree. Since the height of the tree is proportional to *logfan-oud# of data entries),* and the number of disk l/Os to retrieve a data entry is equal to the height (unless some pages are found in the buffer pool), it is clearly important to maximize the fan-out to minimize the height.

An index entry contains a search key value and a page pointer. Hence the size depends primarily on the size of the search key value. If search key values are very long (for instance, the name Devarakonda Venkataramana Sathyanarayana Seshasayee Yellamanchali Murthy, or Donaudampfschifffahrts-kapitansanwiirtersmiitze), not many index entries will fit on a page: Fan-out is low, and the height of the tree is large.

On the other hand, search key values in index entries are used only to direct traffic to the appropriate leaf. When we want to locate data entries with a given search key value, we compare this search key value with the search key values of index entries (on a path from the root to the desired leaf). During the comparison at an index-level node, we want to identify two index entries with search key values kl and k_2 such that the desired search key value k falls between k_1 and $k2$. To accomplish this, we need not store search key values in their entirety in index entries.

For example, suppose we have two adjacent index entries in a node, with search key values 'David Smith' and 'Devarakonda ... ' To discriminate between these two values, it is sufficient to store the abbreviated forms 'Da' and 'De.' More generally, the lneaning of the entry 'David Smith' in the B+ tree is that every value in the subtree pointed to by the pointer to the left of 'David Smith' is less than 'David Smith,' and every value in the subtree pointed to by the pointer to the right of 'David Smith' is (greater than or equal to 'David Smith' and) less than 'Devarakonda ... '

B+ Trees in Real Systems: IBM DB2, Informix, Microsoft SQL Server, Oracle 8, and Sybase ASE all support clustered and unclustered B+ tree indexes, with some differences in how they handle deletions and duplicate key values. In Sybase ASE, depending on the concurrency control schelne being used for the index, the deleted row is removed (with merging if the page occupancy goes below threshold) or simply marked as deleted; a garbage collection scheme is used to recover space in the latter case. In Oracle 8, deletions are handled by marking the row as deleted. To reclaim the space occupied by deleted records, we can rebuild the index online (i.e., while users continue to use the index) or coalesce underfull pages (which does not reduce tree height). Coalesce is in-place, rebuild creates a copy. Informix handles deletions by simply marking records as deleted. DB2 and SQL Server remove deleted records and merge pages when occupancy goes below threshold.

Oracle 8 also allows records from multiple relations to be co-clustered on the same page. The co-clustering can be based on a B+ tree search key or static hashing and up to 32 relations can be stored together.

To ensure such semantics for an entry is preserved, while compressing the entry with key 'David Smith,' we must examine the largest key value in the subtree to the left of 'David Smith' and the smallest key value in the subtree to the right of 'David Smith,' not just the index entries ('Daniel Lee' and 'Devarakonda ...') that are its neighbors. This point is illustrated in Figure 10.21; the value 'Davey Jones' is greater than 'Dav,' and thus, 'David Smith' can be abbreviated only to 'Davi,' not to 'Dav.'

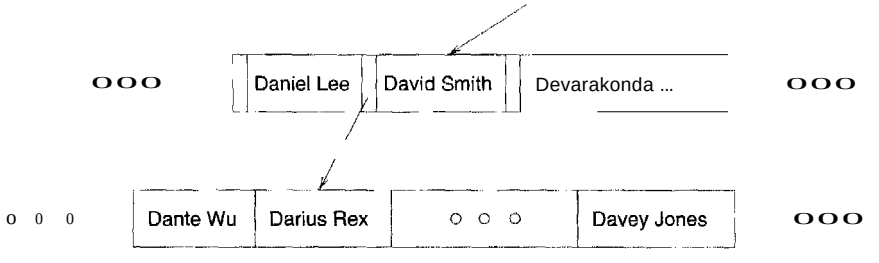


Figure 10.21 Example Illustrating Prefix Key Compression

This technique. called prefix key compression or simply key compression, is supported in many commercial implementations of B+ trees. It can substantially increase the fan-out of a tree. We do not discuss the details of the insertion and deletion algorithms in the presence of key compression.

10.8.2 Bulk-Loading a B+ Tree

Entries are added to a B+ tree in two ways. First, we may have an existing collection of data records with a B+ tree index on it; whenever a record is added to the collection, a corresponding entry must be added to the B+ tree as well. (Of course, a similar comment applies to deletions.) Second, we may have a collection of data records for which we want to create a B+ tree index on some key field(s). In this situation, we can start with an empty tree and insert an entry for each data record, one at a time, using the standard insertion algorithm. However, this approach is likely to be quite expensive because each entry requires us to start from the root and go down to the appropriate leaf page. Even though the index-level pages are likely to stay in the buffer pool between successive requests, the overhead is still considerable.

For this reason many systems provide a *bulk-loading* utility for creating a B+ tree index on an existing collection of data records. The first step is to sort the data entries k^* to be inserted into the (to be created) B+ tree according to the search key k. (If the entries are key-pointer pairs, sorting them does not mean sorting the data records that are pointed to, of course.) We use a running example to illustrate the bulk-loading algorithm. We assume that each data page can hold only two entries, and that each index page can hold two entries and an additional pointer (i.e., the B+ tree is assumed to be of order $d = 1$).

After the data entries have been sorted, we allocate an empty page to serve as the root and insert a pointer to the first page of (sorted) entries into it. We illustrate this process in Figure 10.22, using a sample set of nine sorted pages of data entries.

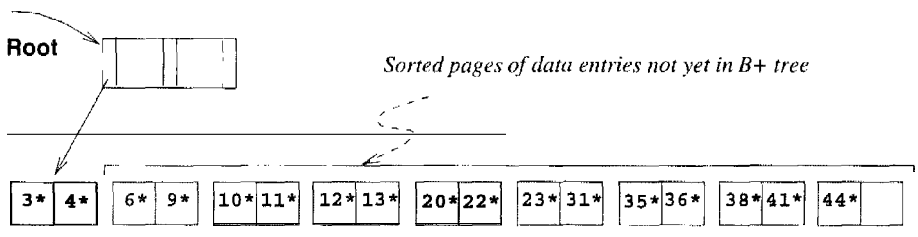


Figure 10.22 Initial Step in B+ Tree Bulk-Loading

We then add one entry to the root page for each page of the sorted data entries. The new entry consists of ⟨*low key value on page, pointer' to page)*. We proceed until the root page is full; see Figure 10.23.

To insert the entry for the next page of data entries, we must split the root and create a new root page. We show this step in Figure 10.24.

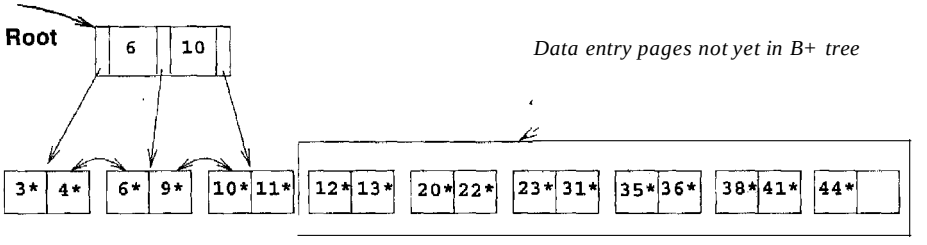


Figure 10.23 Root Page Fills up in B+ Tree Bulk-Loading

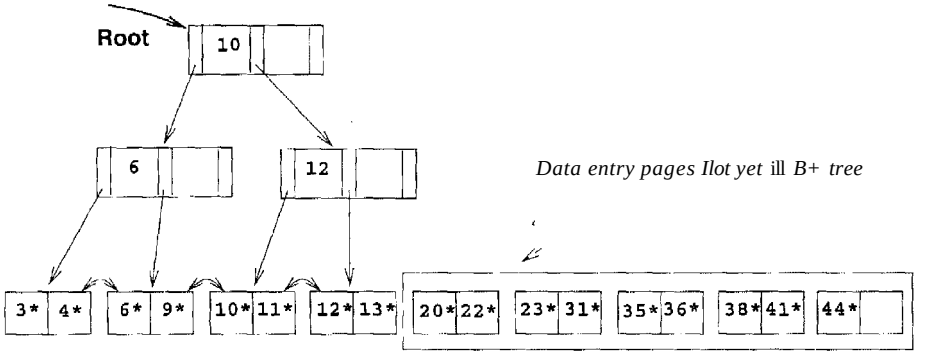


Figure 10.24 Page Split during B+ Tree Bulk-Loading

"We have redistributed the entries evenly between the two children of the root, in anticipation of the fact that the B+ tree is likely to grow. Although it is difficult (!) to illustrate these options when at most two entries fit on a page, we could also have just left all the entries on the old page or filled up some desired fraction of that page (say, 80 percent). These alternatives are simple variants of the basic idea.

To continue with the bulk-loading example, entries for the leaf pages are always inserted into the right-most index page just above the leaf level. 'When the right-most index page above the leaf level fills up, it is split. This action may cause a split of the right-most index page one step closer to the root, as illustrated in Figures 10.25 and 10.26.

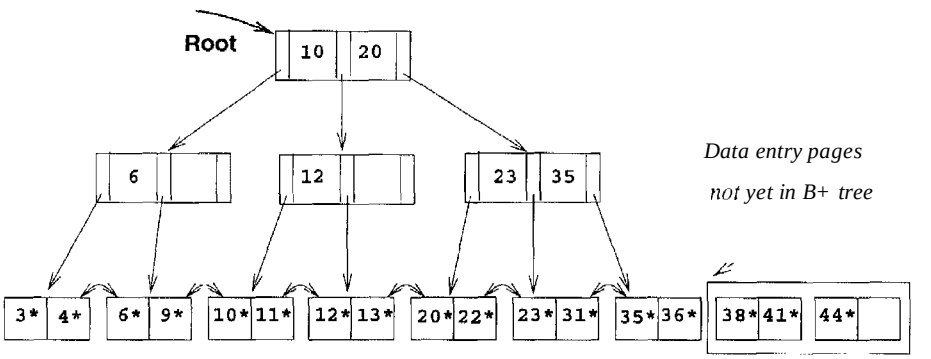


Figure 10.25 Before Adding Entry for Leaf Page Containing 38*

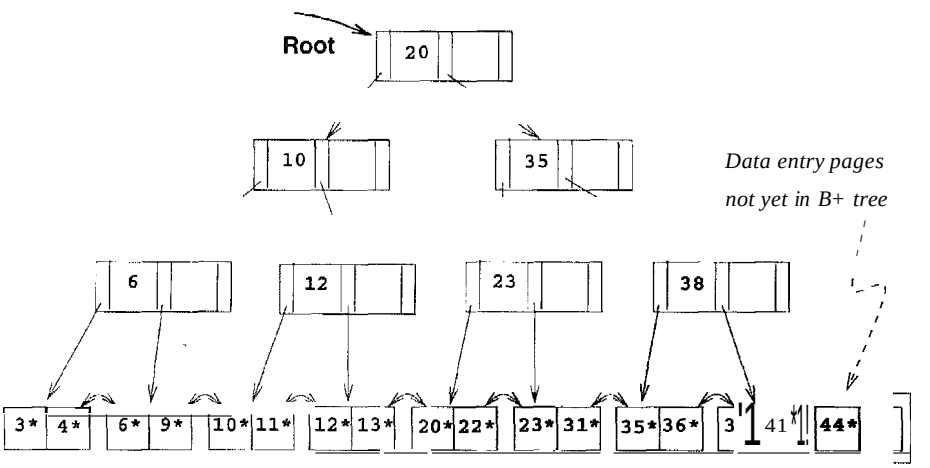


Figure 10.26 After Adding Entry for Leaf Page Containing 38*

Note that splits occur only on the right-most path from the root to the leaf level. We leave the completion of the bulk-loading example as a simple exercise.

Let us consider the cost of creating an index on an existing collection of records. This operation consists of three steps: (1) creating the data entries to insert in the index, (2) sorting the data entries, and (3) building the index from the sorted entries. The first step involves scanning the records and writing out the corresponding data entries; the cost is $(R + E)$ I/Os, where R is the number of pages containing records and E is the number of pages containing data entries. Sorting is discussed in Chapter 13; you will see that the index entries can be generated in sorted order at a cost of about $3E$ I/Os. These entries can then be inserted into the index as they are generated, using the bulk-loading algorithm discussed in this section. The cost of the third step, that is, inserting the entries into the index, is then just the cost of writing out all index pages.

10.8.3 The Order Concept

We presented B+ trees using the parameter *d* to denote minimum occupancy. It is worth noting that the concept of *order* (i.e., the parameter *d),* while useful for teaching B+ tree concepts, must usually be relaxed in practice and replaced by a physical space criterion; for example, that nodes must be kept at least half-full.

One reason for this is that leaf nodes and non-leaf nodes can usually hold different numbers of entries. Recall that B+ tree nodes are disk pages and non-leaf nodes contain only search keys and node pointers, while leaf nodes can contain the actual data records. Obviously, the size of a data record is likely to be quite a bit larger than the size of a search entry, so many more search entries than records fit on a disk page.

A second reason for relaxing the order concept is that the search key may contain a character string field (e.g., the *name* field of Students) whose size varies from record to record; such a search key leads to variable-size data entries and index entries, and the number of entries that will fit on a disk page becomes variable.

Finally, even if the index is built on a fixed-size field, several records may still have the same search key value (e.g., several Students records may have the same *gpa* or *name* value). This situation can also lead to variable-size leaf entries (if we use Alternative (3) for data entries). Because of all these complications, the concept of order is typically replaced by a simple physical criterion (e.g., merge if possible when more than half of the space in the node is unused).

10.8.4 The Effect of Inserts and Deletes on Rids

If the leaf pages contain data records-that is, the B+ tree is a clustered index-then operations such as splits, merges, and redistributions can change rids. Recall that a typical representation for a rid is some combination of (physical) page number and slot number. This scheme allows us to move records within a page if an appropriate page format is chosen but not across pages, as is the case with operations such as splits. So unless rids are chosen to be independent of page numbers, an operation such as split or merge in a clustered B+ tree may require compensating updates to other indexes on the same data.

A similar comment holds for any dynamic clustered index, regardless of whether it is tree-based or hash-based. Of course, the problem does not arise with nonclustered indexes, because only index entries are moved around.

10.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Why are tree-structured indexes good for searches, especially range selections? (Section 10.1)
- Describe how search, insert, and delete operations work in ISAM indexes. Discuss the need for overflow pages, and their potential impact on performance. What kinds of update workloads are ISAM indexes most vulnerable to, and what kinds of workloads do they handle well? (Section 10.2)
- Only leaf pages are affected in updates in ISAM indexes. Discuss the implications for locking and concurrent access. Compare ISAM and B+ trees in this regard. (Section 10.2.1)
- What are the main differences between ISAM and B+ tree indexes? (Section 10.3)
- What is the *order* of a B+ tree? Describe the format of nodes in a B+ tree. Why are nodes at the leaf level linked? (Section 10.3)
- How rnany nodes must be examined for equality search in a B+ tree? How many for a range selection? Compare this with ISAM. (Section 10.4)
- Describe the B+ tree insertion algorithm, and explain how it eliminates overflow pages. Under what conditions can an insert increase the height of the tree? (Section 10.5)
- During deletion, a node might go below the minimum occupancy threshold. How is this handled? Under what conditions could a deletion decrease the height of the tree? (Section 10.6)

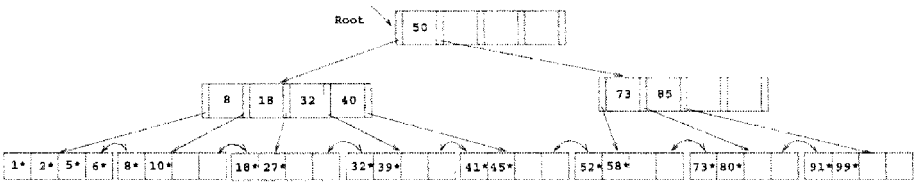


Figure 10.27 Tree for Exercise 10.1

- Why do duplicate search keys require modifications to the implementation of the basic B+ tree operations? (Section 10.7)
- \What is *key compression,* and why is it important? (Section 10.8.1)
- How can a new B+ tree index be efficiently constructed for a set of records? Describe the *bulk-loading* algorithm. (Section 10.8.2)
- Discuss the impact of splits in clustered B+ tree indexes. (Section 10.8.4)

EXERCISES

Exercise 10.1 Consider the B+ tree index of order d = 2 shown in Figure 10.27.

1. Show the tree that would result from inserting a data entry with key 9 into this tree.
2. Show the B+ tree that would result from inserting a data entry with key 3 into the original tree. How many page reads and page writes does the insertion require?
3. Show the B+ tree that would result from deleting the data entry with key 8 from the original tree, assuming that the left sibling is checked for possible redistribution.
4. Show the B+ tree that would result from deleting the data entry with key 8 from the original tree, assuming that the right sibling is checked for possible redistribution.
5. Show the B+ tree that would result from starting with the original tree, inserting a data entry with key 46 and then deleting the data entry with key 52.
6. Show the B+ tree that would result from deleting the data entry with key 91 from the original tree.
7. Show the B+ tree that would result from starting with the original tree, inserting a data entry with key 59, and then deleting the data entry with key 91.
8. Show the B+ tree that \would result from successively deleting the data entries with keys 32, 39, 41, 45, and 73 from the original tree.

Exercise 10.2 Consider the B+ tree index shown in Figure 10.28, which uses Alternative (1) for data entries. Each intermediate node can hold up to five pointers and four key values. Each leaf can hold up to four records, and leaf nodes are doubly linked as usual, although these links are not shown in the figure. Answer the following questions.

1. Name all the tree nodes that mllst be fetched to answer the following query: "Get all records with search key greater than 38."

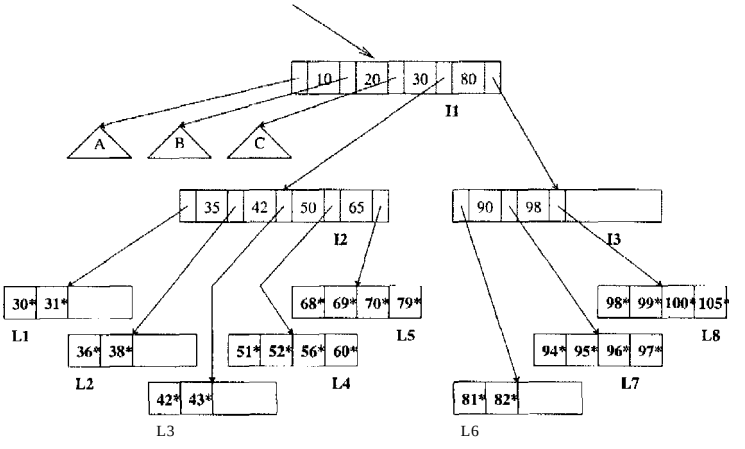


Figure 10.28 Tree for Exercise 10.2

2. Insert a record with search key 109 into the tree.
3. Delete the record with search key 81 from the (original) tree.
4. Name a search key value such that inserting it into the (original) tree would cause an increase in the height of the tree.
5. Note that subtrees A, B, and C are not fully specified. Nonetheless, what can you infer about the contents and the shape of these trees?
6. How would your answers to the preceding questions change if this were an ISAM index?
7. Suppose that this is an ISAM index. What is the minimum number of insertions needed to create a chain of three overflow pages?

Exercise 10.3 Answer the following questions:

1. What is the minimum space utilization for a B+ tree index?
2. What is the minimum space utilization for an ISAM index?
3. If your database system supported both a static and a dynamic tree index (say, ISAM and B+ trees), would you ever consider using the *static* index in preference to the *dynamic* index?

Exercise 10.4 Suppose that a page can contain at most four data values and that all data values are integers. Using only B+ trees of order 2, give examples of each of the following:

1. A B+ tree whose height changes from 2 to 3 when the value 25 is inserted. Show your structure before and after the insertion.
2. A B+ tree in which the deletion of the value 25 leads to a redistribution. Show your structure before and after the deletion.
3. A B+ tree in which the deletion of the value 25 causes a merge of two nodes but without altering the height of the tree.
4. An ISAM structure with four buckets, none of which has an overflow page. Further, every bucket has space for exactly one more entry. Show your structure before and after inserting two additional values, chosen so that an overflow page is created.

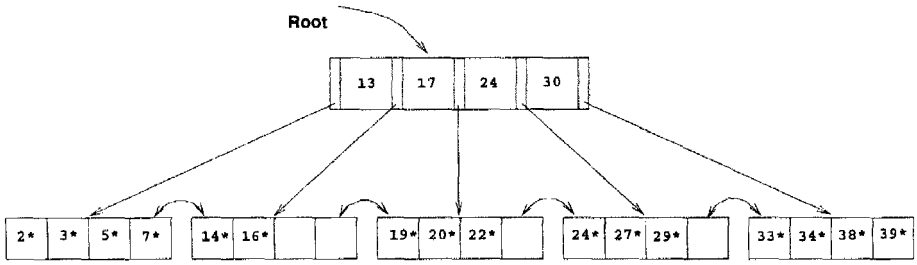


Figure 10.29 Tree for Exercise 10.5

Exercise 10.5 Consider the B+ tree shown in Figure 10.29.

1. Identify a list of five data entries such that:
   (a) Inserting the entries in the order shown and then deleting them in the opposite order (e.g., insert *a*, insert *b*, delete *b*, delete *a)* results in the original tree.
   (b) Inserting the entries in the order shown and then deleting them in the opposite order (e.g., insert *a*, insert *b*, delete *b*, delete *a)* results in a different tree.
2. What is the minimum number of insertions of data entries with distinct keys that will cause the height of the (original) tree to change from its current value (of 1) to 3?
3. Would the minimum number of insertions that will cause the original tree to increase to height 3 change if you were allowed to insert duplicates (multiple data entries with the same key), assuming that overflow pages are not used for handling duplicates?

Exercise 10.6 Answer Exercise 10.5 assuming that the tree is an ISAM tree! (Some of the examples asked for may not exist-if so, explain briefly.)

Exercise 10.7 Suppose that you have a sorted file and want to construct a dense primary B+ tree index on this file.

1. One way to accomplish this task is to scan the file, record by record, inserting each one using the B+ tree insertion procedure. What performance and storage utilization problems are there with this approach?
2. Explain how the bulk-loading algorithm described in the text improves upon this scheme.

Exercise 10.8 Assume that you have just built a dense B+ tree index using Alternative (2) on a heap file containing 20,000 records. The key field for this B+ tree index is a 40-byte string, and it is a candidate key. Pointers (Le., record ids and page ids) are (at most) 10-byte values. The size of one disk page is 1000 bytes. The index was built in a bottom-up fashion using the bulk-loading algorithm, and the nodes at each level were filled up as much as possible.

1. Ho\v many levels does the resulting tree have?
2. For each level of the trec, how many nodes are at that level?
3. How many levels would the resulting tree have if key compression is llsed and it reduces the average size of each key in an entry to 10 bytes?

sid	*name*	*login*	*age*	*gpa*
53831	Maclayall	maclayan@music	11	1.8
53832	Guldu	guldu@music	12	3.8
53666	Jones	jones@cs	18	3.4
53901	Jones	jones@toy	18	3.4
53902	Jones	jones@physics	18	3.4
53903	Jones	jones@english	18	3.4
53904	Jones	jones(ggenetics	18	3.4
53905	Jones	jones@astro	18	3.4
53906	Jones	jones@chem	18	3.4
53902	Jones	jones@sanitation	18	3.8
53688	Smith	smith@ee	19	3.2
53650	Smith	smith@math	19	3.8
54001	Smith	smith@ee	19	3.5
54005	Smith	smith@cs	19	3.8
54009	Smith	smith@astro	19	2.2

Figure 10.30 An Instance of the Students Relation

4. How many levels would the resulting tree have without key compression but with all pages 70 percent full?

Exercise 10.9 The algorithms for insertion and deletion into a B+ tree are presented as recursive algorithms. In the code for *insert*, for instance, a call is made at the parent of a node *N* to insert into (the subtree rooted at) node *N*, and when this call returns, the current node is the parent of *N*. Thus, we do not maintain any 'parent pointers' in nodes of B+ tree. Such pointers are not part of the B+ tree structure for a good reason, as this exercise demonstrates. An alternative approach that uses parent pointers--again, remember that such pointers are *not* part of the standard B+ tree structure!-in each node appears to be simpler:

> Search to the appropriate leaf using the search algorithm; then insert the entry and split if necessary, with splits propagated to parents if necessary (using the parent pointers to find the parents).

Consider this (unsatisfactory) alternative approach:

I. Suppose that an internal node *N* is split into nodes *N* and *N2*. What can you say about the parent pointers in the children of the original node *N?*

2. Suggest two ways of dealing with the inconsistent parent pointers in the children of node *N*.

3. For each of these suggestions, identify a potential (major) disadvantage.

4. What conclusions can you draw from this exercise?

Exercise 10.10 Consider the instance of the Students relation shown in Figure 10.30. Show a B+ tree of order 2 in each of these cases, assuming that duplicates are handled using overflow pages. Clearly indicate what the data entries are (i.e., do not use the *k** convention).

1. A B+ tree index on *age* using Alternative (1) for data entries.
2. A dense B+ tree index on *gpa* using Alternative (2) for data entries. For this question, assume that these tuples are stored in a sorted file in the order shown in the figure: The first tuple is in page 1, slot 1; the second tuple is in page 1, slot 2; and so on. Each page can store up to three data records. You can use *⟨page-id, slot⟩* to identify a tuple.

Exercise 10.11 Suppose that duplicates are handled using the approach without overflow pages discussed in Section 10.7. Describe an algorithm to search for the left-most occurrence of a data entry with search key value K.

Exercise 10.12 Answer Exercise 10.10 assuming that duplicates are handled without using overflow pages, using the alternative approach suggested in Section 9.7.

PROJECT-BASED EXERCISES

Exercise 10.13 Compare the public interfaces for heap files, B+ tree indexes, and linear hashed indexes. What are the similarities and differences? Explain why these similarities and differences exist.

Exercise 10.14 This exercise involves using Minibase to explore the earlier (non-project) exercises further.

1. Create the trees shown in earlier exercises and visualize them using the B+ tree visualizer in Minibase.
2. Verify your answers to exercises that require insertion and deletion of data entries by doing the insertions and deletions in Minibase and looking at the resulting trees using the visualizer.

Exercise 10.15 *(Note to instructors: Additional details must be provided if this exercise* is *assigned; see Appendix 30.)* Implement B+ trees on top of the lower-level code in Minibase.

BIBLIOGRAPHIC NOTES

The original version of the B+ tree was presented by Bayer and McCreight [69]. The B+ tree is described in [442] and [194]. B tree indexes for skewed data distributions are studied in [260]. The VSAM indexing structure is described in [764]. Various tree structures for supporting range queries are surveyed in [79]. An early paper on multiattribute search keys is [498].

References for concurrent access to B+ trees are in the bibliography for Chapter 17.

11

HASH-BASED INDEXING

- What is the intuition behind hash-structured indexes? Why are they especially good for equality searches but useless for range selections?
- What is Extendible Hashing? How does it handle search, insert, and delete?
- What is Linear Hashing? How does it handle search, insert, and delete?
- What are the similarities and differences between Extendible and Linear Hashing?
- **Key concepts:** hash function, bucket, primary and overflow pages, static versus dynamic hash indexes; Extendible Hashing, directory of buckets, splitting a bucket, global and local depth, directory doubling, collisions and overflow pages; Linear Hashing, rounds of splitting, family of hash functions, overflow pages, choice of bucket to split and time to split; relationship between Extendible Hashing's directory and Linear Hashing's family of hash functiolis, need for overflow pages in both schemes in practice, use of a directory for Linear Hashing.

J

Not chaos-like, together crushed and bruised,
But, as the world harmoniously confused:
Where order in variety we see.

Alexander Pope, *Windsor Forest*

In this chapter we consider file organizations that are excellent for equality selections. The basic idea is to use a *hashing function,* which maps values

in a search field into a range of *b'ucket numbers* to find the page on which a desired data entry belongs. We use a simple scheme called *Static Hashing* to introduce the idea. This scheme, like ISAM, suffers from the problem of long overflow chains, which can affect performance. Two solutions to the problem are presented. The *Extendible Hashing* scheme uses a directory to support inserts and deletes efficiently with no overflow pages. The *Linear Hashing* scheme uses a clever policy for creating new buckets and supports inserts and deletes efficiently without the use of a directory. Although overflow pages are used, the length of overflow chains is rarely more than two.

Hash-based indexing techniques cannot support range searches, unfortunately. n'ee-based indexing techniques, discussed in Chapter 10, can support range searches efficiently and are almost as good as hash-based indexing for equality selections. Thus, many commercial systems choose to support only tree-based indexes. Nonetheless, hashing techniques prove to be very useful in implementing relational operations such as joins, as we will see in Chapter 14. In particular, the Index Nested Loops join method generates many equality selection queries, and the difference in cost between a hash-based index and a tree-based index can become significant in this context.

The rest of this chapter is organized as follows. Section 11.1 presents Static Hashing. Like ISAM, its drawback is that performance degrades as the data grows and shrinks. We discuss a dynamic hashing technique, called *Extendible Hashing,* in Section 11.2 and another dynamic technique, called *Linear Hashing,* in Section 11.3. We compare Extendible and Linear Hashing in Section 11.4.

11.1 STATIC HASHING

The Static Hashing scheme is illustrated in Figure 11.1. The pages containing the data can be viewed as a collection of buckets, with one primary page and possibly additional overflow pages per bucket. A file consists of buckets a through $N - 1$, with one primary page per bucket initially. Buckets contain *data entTies,* which can be any of the three alternatives discussed in Chapter 8.

To search for a data entry, we apply a hash function h to identify the bucket to which it belongs and then search this bucket. To speed the search of a bucket, we can maintain data entries in sorted order by search key value; in this chapter, we do not sort entries, and the order of entries within a bucket has no significance. To insert a data entry, we use the hash function to identify the correct bucket and then put the data entry there. If there is no space for this data entry, we allocate a new *overflow* page, put the data entry on this page, and add the page to the overflow chain of the bucket. To delete a data

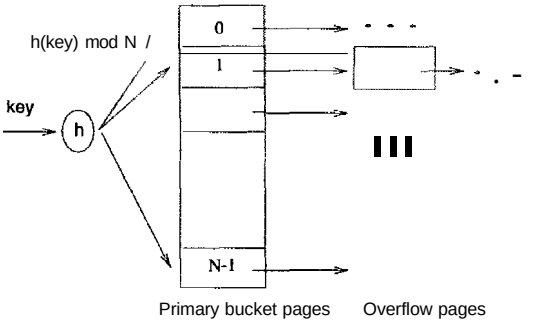


Figure 11.1 Static Hashing

entry, we use the hashing function to identify the correct bucket, locate the data entry by searching the bucket, and then remove it. If this data entry is the last in an overflow page, the overflow page is removed from the overflow chain of the bucket and added to a list of *free pages.*

The hash function is an important component of the hashing approach. It must distribute values in the domain of the search field uniformly over the collection of buckets. If we have N buckets, numbered 0 through $N - 1$, a hash function h of the form $h(value) = (a * value + b)$ works well in practice. (The bucket identified is $h(value) \bmod N$.) The constants a and b can be chosen to 'tune' the hash function.

Since the number of buckets in a Static Hashing file is known when the file is created, the primary pages can be stored on successive disk pages. Hence, a search ideally requires just one disk I/O, and insert and delete operations require two I/Os (read and write the page), although the cost could be higher in the presence of overflow pages. As the file grows, long overflow chains can develop. Since searching a bucket requires us to search (in general) all pages in its overflow chain, it is easy to see how performance can deteriorate. By initially keeping pages 80 percent full, we can avoid overflow pages if the file does not grow too Illuch, but in general the only way to get rid of overflow chains is to create a new file with more buckets.

The main problem with Static Hashing is that the number of buckets is fixed. If a file shrinks greatly, a lot of space is wasted; more important, if a file grows a lot, long overflow chains develop, resulting in poor performance. Therefore, Static Hashing can be compared to the ISAM structure (Section 10.2), which can also develop long overflow chains in case of insertions to the same leaf. Static Hashing also has the same advantages as ISAM with respect to concurrent access (see Section 10.2.1).

One simple alternative to Static Hashing is to periodically 'rehash' the file to restore the ideal situation (no overflow chains, about 80 percent occupancy). However, rehashing takes time and the index cannot be used while rehashing is in progress. Another alternative is to use **dynamic hashing** techniques such as Extendible and Linear Hashing, which deal with inserts and deletes gracefully. We consider these techniques in the rest of this chapter.

11.1.1 Notation and Conventions

In the rest of this chapter, we use the following conventions. As in the previous chapter, record with search key *k,* we denote the index data entry by *k**. For hash-based indexes, the first step in searching for, inserting, or deleting a data entry with search key *k* is to apply a hash function *h* to *k;* we denote this operation by *h(k),* and the value *h(k)* identifies the bucket for the data entry $k*$. Note that two different search keys can have the same hash value.

11.2 EXTENDIBLE HASHING

To understand Extendible Hashing, let us begin by considering a Static Hashing file. If we have to insert a new data entry into a full bucket, we need to add an overflow page. If we do not want to add overflow pages, one solution is to reorganize the file at this point by doubling the number of buckets and redistributing the entries across the new set of buckets. This solution suffers from one major defect--the entire file has to be read, and twice as many pages have to be written to achieve the reorganization. This problem, however, can be overcome by a simple idea: Use a **directory** of pointers to buckets, and double the size of the number of buckets by doubling just the directory and splitting *only* the bucket that overflowed.

To understand the idea, consider the sample file shown in Figure 11.2. The directory consists of an array of size 4, with each element being a pointer to a bucket.. (The *global depth* and *local depth* fields are discussed shortly, ignore them for now.) To locate a data entry, we apply a hash function to the search field and take the last 2 bits of its binary representation to get a number between 0 and 3. The pointer in this array position gives us the desired bucket; we assume that each bucket can hold four data entries. Therefore, to locate a data entry with hash value 5 (binary 101), we look at directory element 01 and follow the pointer to the data page (bucket B in the figure).

To insert a data entry, we search to find the appropriate bucket.. For example, to insert a data entry with hash value 13 (denoted as 13*), we examine directory element 01 and go to the page containing data entries 1*, 5*, and 21*. Since

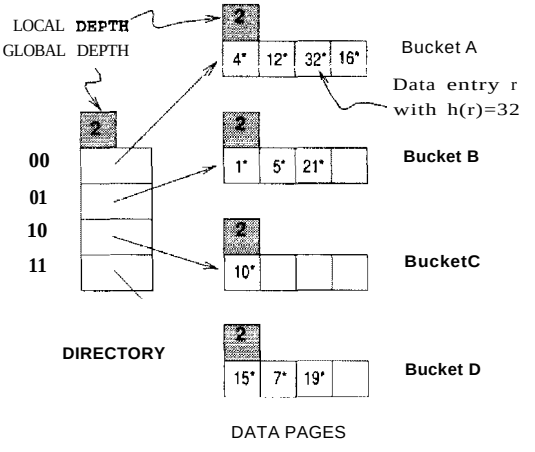


Figure 11.2 Example of an Extendible Hashed File

the page has space for an additional data entry, we are done after we insert the entry (Figure 11.3).

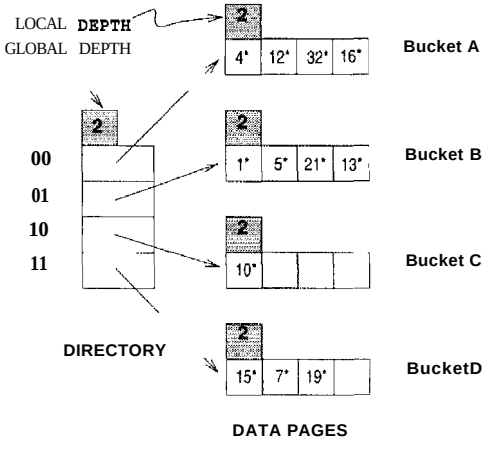


Figure 11.3 After Inserting Entry r with $h(r) = 13$

Next, let us consider insertion of a data entry into a full bucket. The essence of the Extcndible Hashing idea lies in how we deal with this case. Consider the insertion of data entry 20* (binary 10100). Looking at directory clement 00, we arc led to bucket A, which is already full. We 111Ust first **split** the bucket

by allocating a new bucket[1] and redistributing the contents (including the new entry to be inserted) across the old bucket and its 'split image.' To redistribute entries across the old bucket and its split image, we consider the last *three* bits of *h(T);* the last two bits are 00, indicating a data entry that belongs to one of these two buckets, and the third bit discriminates between these buckets. The redistribution of entries is illustrated in Figure 11.4.

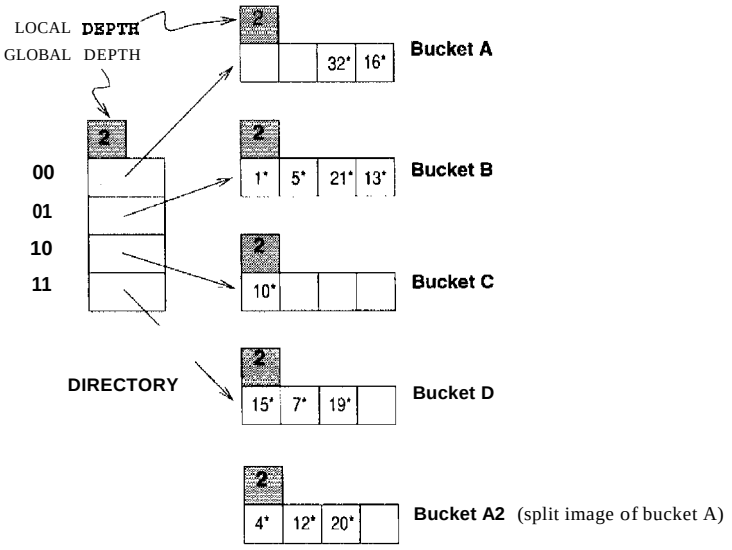


Figure 11.4 While Inserting Entry r with *h(r}=20*

Note a problem that we must now resolve—we need three bits to discriminate between two of our data pages (A and A2), but the directory has only enough slots to store all two-bit patterns. The solution is to *double the directory.* Elements that differ only in the third bit from the end are said to 'correspond': *COT-r'esponding elements* of the directory point to the same bucket with the exception of the elements corresponding to the split bucket. In our example, bucket a was split; so, new directory element 000 points to one of the split versions and new element 100 points to the other. The sample file after completing all steps in the insertion of 20* is shown in Figure 11.5.

Therefore, doubling the file requires allocating a new bucket page, writing both this page and the old bucket page that is being split, and doubling the directory array. The directory is likely to be much smaller than the file itself because each element is just a page-id, and can be doubled by simply copying it over

[1]Since there are 'no overflow pages in Extendible Hashing, a bucket can be thought of as a single page.

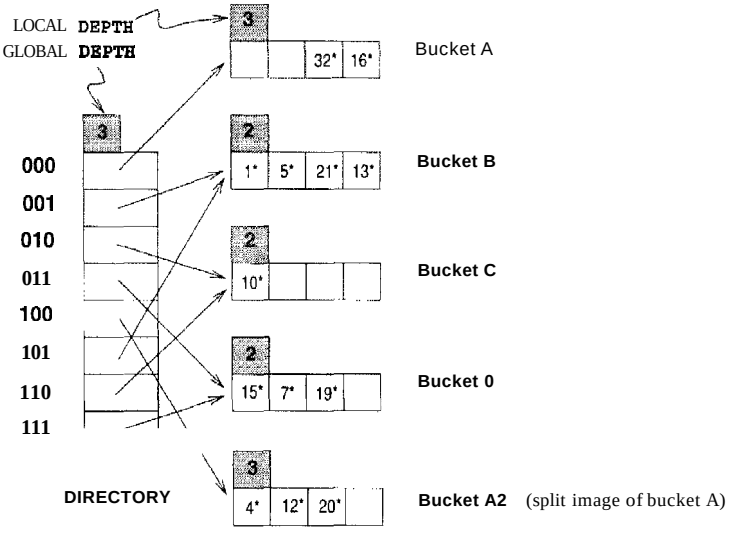


Figure 11.5 After Inserting Entry r with $h(r) = 20$

(and adjusting the elements for the split buckets). The cost of doubling is now quite acceptable.

We observe that the basic technique used in Extendible Hashing is to treat the result of applying a hash function h as a binary number and interpret the last d bits, where d depends on the size of the directory, as an offset into the directory. In our example, d is originally 2 because we only have four buckets; after the split, d becomes 3 because we now have eight buckets. A corollary is that, when distributing entries across a bucket and its split image, we should do so on the basis of the dth bit. (Note how entries are redistributed in our example; see Figure 11.5.) The number d, called the **global depth** of the hashed file, is kept as part of the header of the file. It is used every time we need to locate a data entry.

An important point that arises is whether splitting a bucket necessitates a directory doubling. Consider our example, as shown in Figure 11.5. If we now insert 9*, it belongs in bucket B; this bucket is already full. We can deal with this situation by splitting the bucket and using directory elements 001 and 10] to point to the bucket and its split image, as shown in Figure 11.6.

Hence, a bucket split does not necessarily require a directory doubling. However, if either bucket A or A2 grows full and an insert then forces a bucket split, we are forced to double the directory again.

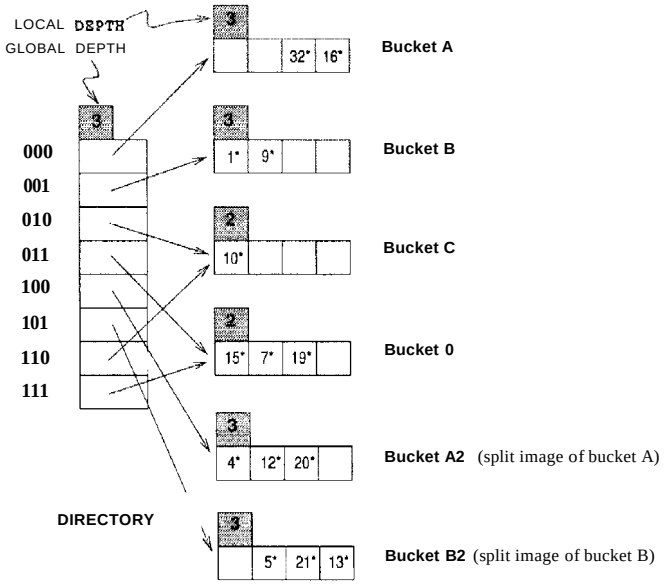


Figure 11.6 After Inserting Entry r with $h(r) = 9$

To differentiate between these cases and determine whether a directory doubling is needed, we maintain a **local depth** for each bucket. If a bucket whose local depth is equal to the global depth is split, the directory must be doubled. Going back to the example, when we inserted 9* into the index shown in Figure 11.5, it belonged to bucket B with local depth 2, whereas the global depth was 3. Even though the bucket was split, the directory did not have to be doubled. Buckets A and A2, on the other hand, have local depth equal to the global depth, and, if they grow full and are split, the directory must then be doubled.

Initially, all local depths are equal to the global depth (which is the number of bits needed to express the total number of buckets). We increment the global depth by 1 each time the directory doubles, of course. Also, whenever a bucket is split (whether or not the split leads to a directory doubling), we increment by 1 the local depth of the split bucket and assign this same (incremented) local depth to its (newly created) split image. Intuitively, if a bucket has local depth l, the hash values of data entries in it agree on the last l bits; further, no data entry in any other bucket of the file has a hash value with the same last l bits. A total of 2^{d-l} directory elernents point to a bucket with local depth l; if $d = l$, exactly one directory element points to the bucket and splitting such a bucket requires directory doubling.

A final point to note is that we can also use the first *d* bits (the *most significant* bits) instead of the last *d (least significant* bits), but in practice the *last d* bits are used. The reason is that a directory can then be doubled simply by copying it.

In summary, a data entry can be located by computing its hash value, taking the last *d* bits, and looking in the bucket pointed to by this directory element. For inserts, the data entry is placed in the bucket to which it belongs and the bucket is split if necessary to make space. A bucket split leads to an increase in the local depth and, if the local depth becomes greater than the global depth as a result, to a directory doubling (and an increase in the global depth) as well.

For deletes, the data entry is located and removed. If the delete leaves the bucket empty, it can be merged with its split image, although this step is often omitted in practice. Merging buckets decreases the local depth. If each directory element points to the same bucket as its split image (i.e., 0 and 2^{d-1} point to the same bucket, namely, A; 1 and $2^{d-1} + 1$ point to the same bucket, namely, B, which mayor may not be identical to A; etc.), we can halve the directory and reduce the global depth, although this step is not necessary for correctness.

The insertion examples can be worked out backwards as examples of deletion. (Start with the structure shown after an insertion and delete the inserted element. In each case the original structure should be the result.)

If the directory fits in memory, an equality selection can be answered in a single disk access, as for Static Hashing (in the absence of overflow pages), but otherwise, two disk I/Os are needed. As a typical example, a 100MB file with 100 bytes per data entry and a page size of 4KB contains 1 million data entries and only about 25,000 elements in the directory. (Each page/bucket contains roughly 40 data entries, and we have one directory element per bucket.) Thus, although equality selections can be twice as slow as for Static Hashing files, chances are high that the directory will fit in memory and performance is the same as for Static Hashing files.

On the other hand, the directory grows in spurts and can become large for *skewed data distributions* (where our assumption that data pages contain roughly equal numbers of data entries is not valid). In the context of hashed files, in a skewed data distribution the distribution of *hash values of search field values* (rather than the distribution of search field values themselves) is skewed (very 'bursty' or nonuniform). Even if the distribution of search values is skewed, the choice of a good hashing function typically yields a fairly uniform distribution of hash va.lues; skew is therefore not a problem in practice.

F\lrther, collisions, or data entries with the same hash value, cause a problem and must be handled specially: \Vhen more data entries than \vill fit on a page have the same hash value, we need overflow pages.

11.3 LINEAR HASHING

Linear Hashing is a dynamic hashing technique, like Extendible Hashing, adjusting gracefully to inserts and deletes. In contrast to Extendible Hashing, it does not require a directory, deals naturally with collisions, and offers a lot of flexibility with respect to the timing of bucket splits (allowing us to trade off slightly greater overflow chains for higher average space utilization). If the data distribution is very skewed, however, overflow chains could cause Linear Hashing performance to be worse than that of Extendible Hashing.

The scheme utilizes a *family* of hash functions h_a, hI, *h2,* ..., with the property that each function's range is twice that of its predecessor. That is, if *hi* maps a data entry into one of M buckets, h_i+I maps a data entry into one of $2M$ buckets. Such a family is typically obtained by choosing a hash function hand an initial number N ofbuckets,[2] and defining *hi(value) = h(value) mod* $(2^i N)$. If N is chosen to be a power of 2, then we apply h and look at the last d_i bits; d_o is the number of bits needed to represent N, and $d_i = d_a + i$. Typically we choose h to be a function that maps a data entry to some integer. Suppose that we set the initial number N of buckets to be 32. In this case d_o is 5, and h_a is therefore $h \bmod 32$, that is, a number in the range 0 to 31. The value of d_1 is $d_o + 1 = 6$, and hI is $h \bmod (2 * 32)$, that is, a number in the range 0 to 63. Then h_2 yields a number in the range 0 to 127, and so oII.

The idea is best understood in terms of **rounds** of splitting. During round number *Level,* only hash functions h_{Level} and $h_{Level+1}$ are in use. The buckets in the file at the beginning of the round are split, one by one from the first to the last bucket, thereby doubling the number of buckets. At any given point within a round, therefore, we have buckets that have been split, buckets that are yet to be split, and buckets created by splits in this round, as illustrated in Figure 11.7.

Consider how we search for a data entry with a given search key value. \Ve apply hash function h_{Level}, and if this leads us to one of the unsplit buckets, we simply look there. If it leads us to one of the split buckets, the entry may be there or it may have been moved to the new bucket created earlier in this round by splitting this bucket; to determine which of the two buckets contains the entry, we apply $h_{Level+1}$.

[2] Note that 0 to IV - 1 is *not* the range of fl.!

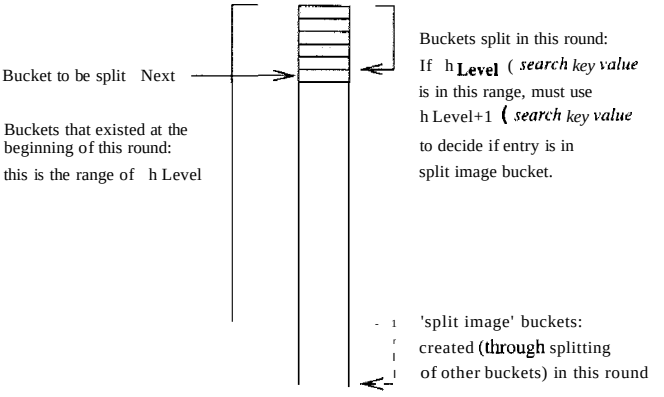


Figure 11.7 Buckets during a Round in Linear Hashing

Unlike Extendible Hashing, when an insert triggers a split, the bucket into which the data entry is inserted is not necessarily the bucket that is split. An overflow page is added to store the newly inserted data entry (which triggered the split), as in Static Hashing. However, since the bucket to split is chosen in round-robin fashion, eventually all buckets are split, thereby redistributing the data entries in overflow chains before the chains get to be more than one or two pages long.

We now describe Linear Hashing in more detail. A counter *Level* is used to indicate the current round number and is initialized to 0. The bucket to split is denoted by *Next* and is initially bucket 0 (the first bucket). We denote the number of buckets in the file at the beginning of round *Level* by N_{Level}. We can easily verify that $N_{Level} = N * 2^{Level}$. Let the number of buckets at the beginning of round 0, denoted by N_0, be N. We show a small linear hashed file in Figure 11.8. Each bucket can hold four data entries, and the file initially contains four buckets, as shown in the figure.

We have considerable flexibility in how to trigger a split, thanks to the use of overflow pages. We can split whenever a new overflow page is added, or we can impose additional conditions based all conditions such as space utilization. For our examples, a split is 'triggered' when inserting a new data entry causes the creation of an Qverftow page.

\Vhenever a split is triggered the *Next* bucket is split, and hash function $h_{Level+1}$ redistributes entries between this bucket (say bucket number b) and its split image; the split image is therefore bucket number $b + N_{Level}$. After splitting a bucket, the value of *Next* is incremented by 1. In the example file, insertion of

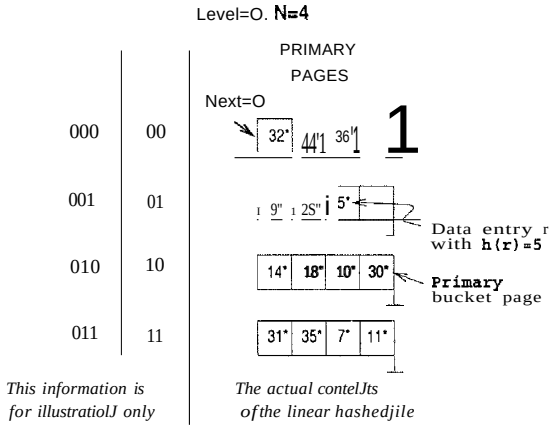


Figure 11.8 Example of a Linear Hashed File

data entry 43* triggers a split. The file after completing the insertion is shown in Figure 11.9.

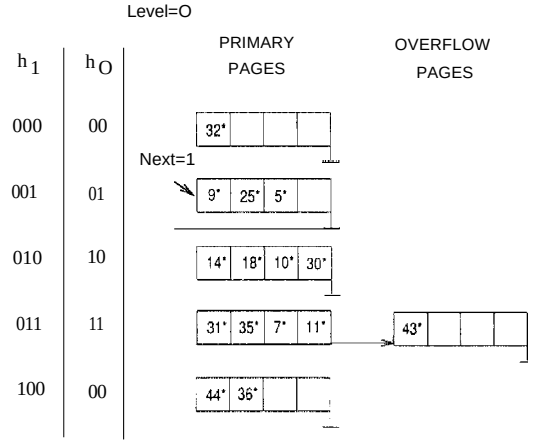


Figure 11.9 After Inserting Record r with $h(T) = 43$

At any time in .the middle of a round *Level,* all buckets above bucket *Next* have been split, and the file contains buckets that are their split images, as illustrated in Figure 11.7. Buckets *Next* through *NLevcl* have not yet been split. **If** we use *hLevel* on a data entry and obtain a number *b* in the range *Next* through *NLevel,* the data entry belongs to bucket *b.* For example, $h_O(18)$ is 2 (binary 10); since this value is between the current values of *Ne:r:t* (= 1) and N_1 (= 4), this bucket has not been split. However, if we obtain a number *b* in the range 0 through

Next, the data entry may be in this bucket or in its split image (which is bucket number $b + N_{Level}$); we have to use $h_{Level+1}$ to determine to which of these two buckets the data entry belongs. In other words, we have to look at one more bit of the data entry's hash value. For example, $h_0(32)$ and $h_0(44)$ are both a (binary 00). Since *Next* is currently equal to 1, which indicates a bucket that has been split, we have to apply hI' We have hI(32) = 0 (binary 000) and $h_1(44)$ = 4 (binary 100). Therefore, 32 belongs in bucket A and 44 belongs in its split image, bucket A2.

Not all insertions trigger a split, of course. If we insert 37* into the file shown in Figure 11.9, the appropriate bucket has space for the new data entry. The file after the insertion is shown in Figure 11.10.

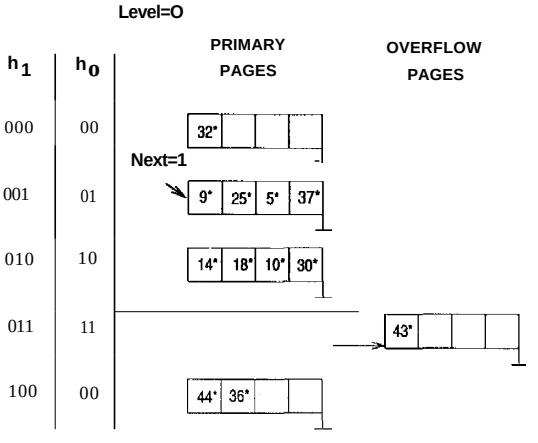


Figure 11.10 After Inserting Record r with $h(r)$ = 37

Sometimes the bucket pointed to by *Next* (the current candidate for splitting) is full, and a new data entry should be inserted in this bucket. In this case, a split is triggered, of course, but we do not need a new overflow bucket. This situation is illustrated by inserting 29* into the file shown in Figure 11.10. The result is shown in Figure 11.11.

When *Next* is equal to N_{Level} - 1 and a split is triggered, we split the last of the buckets present in the file at the beginning of round *Level.* The number of buckets after the split is twice the number at the beginning of the round, and we start a new round with *Level* incremented by 1 and *Next* reset to 0. Incrementing *Level* amounts to doubling the effective range into which keys are hashed. Consider the example file in Figure 11.12, which was obtained from the file of Figure 11.11 by inserting 22*, 66*, and 34*. (The reader is encouraged to try to work out the details of these insertions.) Inserting 50* causes a split that

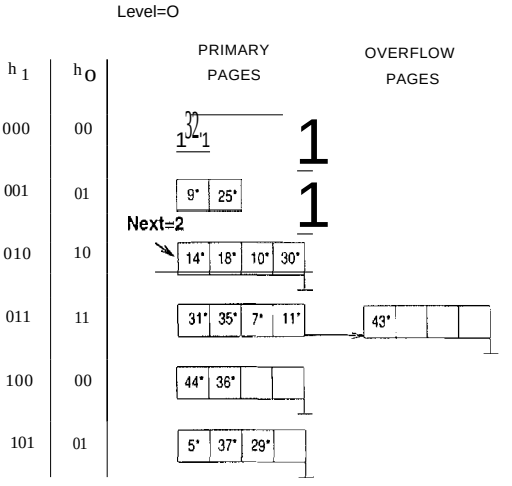


Figure 11.11 After Inserting Record r with $h(r^*) = 29$

leads to incrementing *Level,* as discussed previously; the file after this insertion is shown in Figure 11.13.

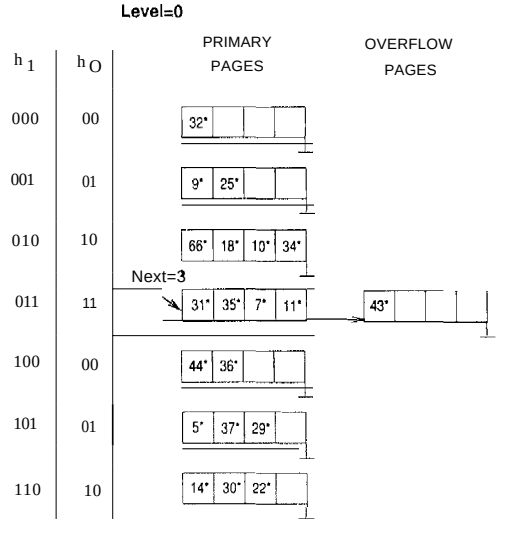


Figure 11.12 After Inserting Records with $h(r) = 22,66,and34$

In summary, an equality selection costs just one disk I/O unless the bucket has overflow pages; in practice, the cost on average is about 1.2 disk accesses for

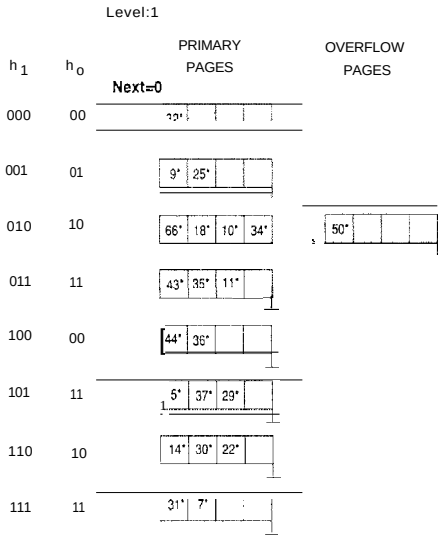


Figure 11.13 After Inserting Record r with $h(r) = 50$

reasonably uniform data distributions. (The cost can be considerably worse--linear in the number of data entries in the file----if the distribution is very skewed. The space utilization is also very poor with skewed data distributions.) Inserts require reading and writing a single page, unless a split is triggered.

We not discuss deletion in detail, but it is essentially the inverse of insertion. If the last bucket in the file is empty, it can be removed and *Next* can be decremented. (If *Next* is 0 and the last bucket becomes empty, *Next* is made to point to bucket $(M/2) - 1$, where *!vI* is the current number of buckets, *Level* is decremented, and the empty bucket is removed.) If we wish, we can combine the last bucket with its split image even when it is not empty, using some criterion to trigger this merging in essentially the same way. The criterion is typically based on the occupancy of the file, and merging can be done to improve space utilization.

11.4 EXTENDIBLE VS. LINEAR HASHING

To understand the relationship between Linear Hashing and Extendible Hashing, imagine that we also have a directory in Linear Hashing with elements 0 to $N - 1$. The first split is at bucket 0, and so we add directory element N. In principle, we may imagine that the entire directory has been doubled at this point; however, because element 1 is the same as element $N + 1$, elernent 2 is

the same as element $N + 2$, and so on, we can avoid the actual copying for the rest of the directory. The second split occurs at bucket 1; now directory element $N + 1$ becomes significant and is added. At the end of the round, all the original N buckets are split, and the directory is doubled in size (because all elements point to distinct buckets).

We observe that the choice of hashing functions is actually very similar to what goes on in Extendible Hashing---in effect, moving from h_i to h_{i+1} in Linear Hashing corresponds to doubling the directory in Extendible Hashing. Both operations double the effective range into which key values are hashed; but whereas the directory is doubled in a single step of Extendible Hashing, moving from h_i to h_{i+1}, along with a corresponding doubling in the number of buckets, occurs gradually over the course of a round in Linear Hashing. The new idea behind Linear Hashing is that a directory can be avoided by a clever choice of the bucket to split. On the other hand, by always splitting the appropriate bucket, Extendible Hashing may lead to a reduced number of splits and higher bucket occupancy.

The directory analogy is useful for understanding the ideas behind Extendible and Linear Hashing. However, the directory structure can be avoided for Linear Hashing (but not for Extendible Hashing) by allocating primary bucket pages consecutively, which would allow us to locate the page for bucket i by a simple offset calculation. For uniform distributions, this implementation of Linear Hashing has a lower average cost for equality selections (because the directory level is eliminated). For skewed distributions, this implementation could result in any empty or nearly empty buckets, each of which is allocated at least one page, leading to poor performance relative to Extendible Hashing, which is likely to have higher bucket occupancy.

A different implementation of Linear Hashing, in which a directory is actually maintained, offers the flexibility of not allocating one page per bucket; *null* directory elements can be used as in Extendible Hashing. However, this implementation introduces the overhead of a directory level and could prove costly for large, uniformly distributed files. (Also, although this implementation alleviates the potential problem of low bucket occupancy by not allocating pages for empty buckets, it is not a complete solution because we can still have many pages with very few entries.)

11.5 REVIEW **QUESTIONS**

Answers to the review questions can be found in the listed sections.

- How does a hash-based index handle an equality query? Discuss the use of the hash function in identifying a bucket to search. Given a bucket number, explain how the record is located on disk.
- Explain how insert and delete operations are handled in a static hash index. Discuss how overflow pages are used, and their impact on performance. How many disk I/Os does an equality search require, in the absence of overflow chains? What kinds of workload does a static hash index handle well, and when it is especially poor? **(Section 11.1)**
- How does Extendible Hashing use a directory of buckets? How does Extendible Hashing handle an equality query? How does it handle insert and delete operations? Discuss the *global depth* of the index and *local depth* of a bucket in your answer. Under what conditions can the directory can get large? **(Section 11.2)**
- What are *collisions?* Why do we need overflow pages to handle them? **(Section 11.2)**
- How does *Linear Hashing* avoid a directory? Discuss the round-robin splitting of buckets. Explain how the split bucket is chosen, and what triggers a split. Explain the role of the family of hash functions, and the role of the *Level* and *Next* counters. When does a round of splitting end? **(Section 11.3)**
- Discuss the relationship between Extendible and Linear Hashing. What are their relative merits? Consider space utilization for skewed distributions, the use of overflow pages to handle collisions in Extendible Hashing, and the use of a directory in Linear Hashing. **(Section 11.4)**

EXERCISES

Exercise **11.1** Consider the Extendible Hashing index shown in Figure 11.14. Answer the following questions about this index:

1. What can you say about the last entry that was inserted into the index?
2. What can you say about the last entry that was inserted into the index if you know that there have been no deletions from this index so far?
3. Suppose you are told that there have been no deletions from this index so far. What can you say about the last entry whose insertion into the index caused a split?
4. Show the index after inserting an entry with hash value 68.
5. Show the original index after inserting entries with hash values 17 and 69.
6. Show the original index after deleting the entry with hash value 21. (Assume that the full deletion algorithm is used.)
7. Show the original index after deleting the entry with hash value 10. Is a merge triggered by this deletion? If not, explain why. (Assume that the full deletion algorithm is used.)

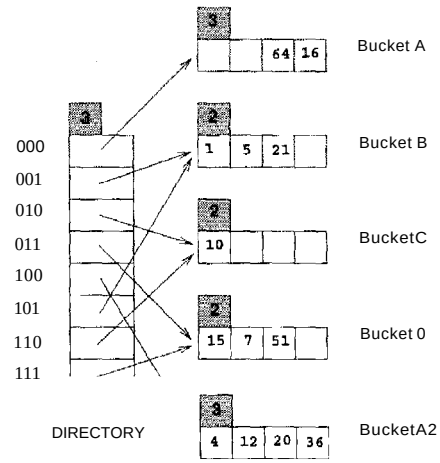


Figure 11.14 Figure for Exercise 11.1

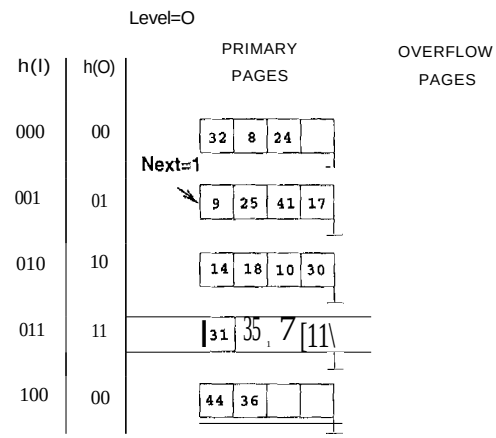


Figure 11.15 Figure for Exercise 11.2

Exercise 11.2 Consider the Linear Hashing index shown in Figure 11.15. Assume that we split whenever an overflow page is created. Answer the following questions about this index:

1. What can you say about the last entry that was inserted into the index?
2. What can you say about the last entry that was inserted into the index if you know that there llave been no deletions from this index so far?
3. Suppose you know that there have been no deletions from this index so far. What can you say about the last entry whose insertion into the index caused a split?
4. Show the index after inserting an entry with hash value 4.

5. Show the original index after inserting an entry with hash value 15.
6. Show the original index after deleting the entries with hash values 36 and 44. (Assume that the full deletion algorithm is used.)
7. Find a list of entries whose insertion into the original index would lead to a bucket with two overflow pages. Use as few entries as possible to accomplish this. What is the maximum number of entries that can be inserted into this bucket before a split occurs that reduces the length of this overflow chain?

Exercise 11.3 Answer the following questions about Extendible Hashing:

1. Explain why local depth and global depth are needed.
2. After an insertion that causes the directory size to double, how many buckets have exactly one directory entry pointing to them? If an entry is then deleted from one of these buckets, what happens to the directory size? Explain your answers briefly.
3. Does Extendible I-lashing guarantee at most one disk access to retrieve a record with a given key value?
4. If the hash function distributes data entries over the space of bucket numbers in a very skewed (non-uniform) way, what can you say about the size of the directory? What can you say about the space utilization in data pages (i.e., non-directory pages)?
5. Does doubling the directory require us to examine all buckets with local depth equal to global depth?
6. Why is handling duplicate key values in Extendible Hashing harder than in ISAM?

Exercise 11.4 Answer the following questions about Linear Hashing:

1. How does Linear Hashing provide an average-case search cost of only slightly more than one disk I/O, given that overflow buckets are part of its data structure?
2. Does Linear Hashing guarantee at most one disk access to retrieve a record with a given key value?
3. If a Linear Hashing index using Alternative (1) for data entries contains N records, with P records per page and an average storage utilization of 80 percent, what is the worst-case cost for an equality search? Under what conditions would this cost be the actual search cost?
4. If the hash function distributes data entries over the space of bucket numbers in a very skewed (non-uniform) way, what can you say about thc space utilization in data pages?

Exercise 11.5 Give an example of when you would use each element (A or B) for each of the following 'A versus B' pairs:

1. A hashed index using Alternative (1) versus heap file organization.
2. Extendible Hashing versus Linear Hashing.
3. Static Hashing versus Linear Hashing.
4. Static Hashing versus ISAIVI.
5. Linear Hashing versus B+ trees.

Exercise 11.6 Give examples of the following:

1. A Linear Hashing index and an Extendible Hashing index with the same data entries, such that the Linear Hashing index has more pages.

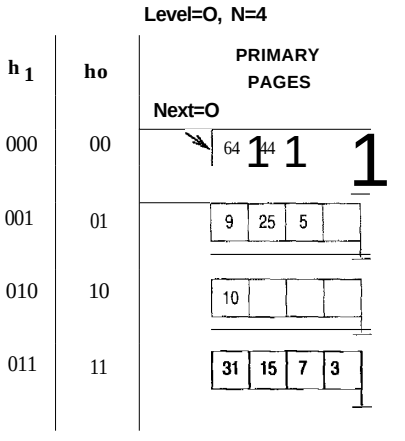


Figure **11.16** Figure for Exercise 11.9

2. A Linear Hashing index and an Extendible Hashing index with the same data entries, such that the Extendible Hashing index has more pages.

Exercise **11.7** Consider a relation R(*a*, *b*, c, *d*) containing 1 million records, where each page of the relation holds 10 records. R is organized as a heap file with unclustered indexes, and the records in R are randomly ordered. Assume that attribute *a* is a candidate key for R, with values lying in the range 0 to 999,999. For each of the following queries, name the approach that would most likely require the fewest I/Os for processing the query. The approaches to consider follow:

- Scanning through the whole heap file for R.
- Using a B+ tree index on attribute *R.a*.
- Using a hash index on attribute *R.a*.

The queries are:

1. Find all R tuples.
2. Find all R tuples such that $a < 50$.
3. Find all R tuples such that $a = 50$.
4. Find all R tuples such that $a > 50$ and $a < 100$.

Exercise 11.8 How would your answers to Exercise 11.7 change if *a* is not a candidate key for R? How would thcy change if we assume that records in R are sorted on *a*?

Exercise **11.9** Consider the snapshot of the Linear Hashing index shown in Figure 11.16. Assume that a bucket split occurs whcnever an overflow page is created.

1. What is the *maximum* number of data entries that call be inserted (given the best possible distribution of keys) before you have to split a bucket? Explain very briefly.
2. Show the file after inserting a *single* record whose insertion causes a bucket split.

3. (a) What is the *minimum* number of record insertions that will cause a split of all four buckets? Explain very briefly.
   (b) What is the value of *Next* after making these insertions?
   (c) What can you say about the number of pages in the fourth bucket shown after this series of record insertions?

Exercise **11.10** Consider the data entries in the Linear Hashing index for Exercise 11.9.

1. Show an Extendible Hashing index with the same data entries.
2. Answer the questions in Exercise 11.9 with respect to this index.

Exercise **11.11** In answering the following questions, assume that the full deletion algorithm is used. Assume that merging is done when a bucket becomes empty.

1. Give an example of Extendible Hashing where deleting an entry reduces global depth.
2. Give an example of Linear Hashing in which deleting an entry decrements *Next* but leaves *Level* unchanged. Show the file before and after the deletion.
3. Give an example of Linear Hashing in which deleting an entry decrements *Level*. Show the file before and after the deletion.
4. Give an example of Extendible Hashing and a list of entries e1, *e2, e3* such that inserting the entries in order leads to three splits and deleting them in the reverse order yields the original index. If such an example does not exist, explain.
5. Give an example of a Linear Hashing index and a list of entries e1, *e2, e3* such that inserting the entries in order leads to three splits and deleting them in the reverse order yields the original index. If such an example does not exist, explain.

PROJECT-BASED EXERCISES

Exercise **11.12** *(Note to instructors: Additional details must be provided if this question is assigned. See Appendix 30.)* Implement Linear Hashing or Extendible Hashing in Minibase.

BIBLIOGRAPHIC NOTES

Hashing is discussed in detail in [442]. Extendible Hashing is proposed in [256]. Litwin proposed Linear Hashing in [483]. A generalization of Linear Hashing for distributed environments is described in [487]. There has been extensive research into hash-based indexing techniques. Larson describes two variations of Linear Hashing in [469] and [470]. Ramakrishna presents an analysis of hashing techniques in [607]. Hash functions that do not produce bucket overflows are studied in [608]. Order-preserving hashing techniques are discussed in [484] and [308]. Partitioned-hashing, in which each field is hashed to obtain some bits of the bucket address, extends hashing for the case of queries in which equality conditions are specified only for some of the key fields. This approach was proposed by Rivest [628] and is discussed in [747]; a further development is described in [616].

PARTN

QUERY EVALUATION

12

OVERVIEW OF QUERY EVALUATION

- What descriptive information does a DBMS store in its catalog?
- What alternatives are considered for retrieving rows from a table?
- Why does a DBMS implement several algorithms for each algebra operation? What factors affect the relative performance of different algorithms?
- What are query evaluation plans and how are they represented?
- Why is it important to find a good evaluation plan for a query? How is this done in a relational DBMS?
- Key concepts: catalog, system statistics; fundamental techniques, indexing, iteration, and partitioning; access paths, matching indexes and selection conditions; selection operator, indexes versus scans, impact of clustering; projection operator, duplicate elimination; join operator, index nested-loops join, sort-merge join; query evaluation plan; materialization vs. pipelinining; iterator interface; query optimization, algebra equivalences, plan enumeration; cost estimation

This very remarkable man, commends a most practical plan:
You can do what you want, if you don't think you can't,
So c1on't think you can't if you can.

—Charles Inge

In this chapter, we present an overview of how queries are evaluated in a relational DBMS. We begin with a discussion of how a DBMS describes the data

that it manages, including tables and indexes, in Section 12.1. This descriptive data, or metadata, stored in special tables called the system catalogs, is used to find the best way to evaluate a query.

SQL queries are translated into an extended form of relational algebra, and query evaluation plans are represented as trees of relational operators, along with labels that identify the algorithm to use at each node. Thus, relational operators serve as building blocks for evaluating queries, and the implementation of these operators is carefully optimized for good performance. We introduce operator evaluation in Section 12.2 and describe evaluation algorithms for various operators in Section 12.3.

In general, queries are composed of several operators, and the algorithms for individual operators can be combined in many ways to evaluate a query. The process of finding a good evaluation plan is called *query optimization*. We introduce query optimization in Section 12.4. The basic task in query optimization, which is to consider several alternative evaluation plans for a query, is motivated through examples in Section 12.5. In Section 12.6, we describe the space of plans considered by a typical relational optimizer.

The ideas are presented in sufficient detail to allow readers to understand how current database systems evaluate typical queries. This chapter provides the necessary background in query evaluation for the discussion of physical database design and tuning in Chapter 20. Relational operator implementation and query optimization are discussed further in Chapters 13, 14, and 15; this in-depth coverage describes how current systems are implemented.

We consider a number of example queries using the following schema:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)
Reserves(sid: integer, *bid:* integer, *day:* dates, *marne:* string)

We aSSUlne that each tuple of Reserves is 40 bytes long, that a page can hold 100 Reserves tuples, and that we have 1000 pages of such tuples. Similarly, we assume that each tuple of Sailors is 50 bytes long, that a page can hold 80 Sailors tuples, and that we have 500 pages of such tuples.

12.1 THE SYSTEM CATALOG

We can store a table using one of several alternative file structures, and we can create one or more indexes -each stored as a file 011 every tal)le. Conversely, in a relational DBMS, every file contains either the tuples in a table or the

entries in an index. The collection of files corresponding to users' tables and indexes represents the *data* in the database.

A relational DBMS maintains information about every table and index that it contains. The descriptive information is itself stored in a collection of special tables called the catalog tables. An example of a catalog table is shown in Figure 12.1. The catalog tables are also called the data dictionary, the system catalog, or simply the *catalog.*

12.1.1 Information in the Catalog

Let us consider what is stored in the system catalog. At a minimum, we "have system-wide information, such as the size of the buffer pool and the page size, and the following information about individual tables, indexes, and views:

- For each table:
  - Its *table name,* the *file name* (or some identifier), and the *file structure* (e.g., heap file) of the file in which it is stored.
  - The *attribute name* and *type* of each of its attributes.
  - The *index name* of each index on the table.
  - The *integrity constmints* (e.g., primary key and foreign key constraints) on the table.
- For each index:
  - The *index name* and the *structure* (e.g., B+ tree) of the index.
  - The *search key* attributes.
- For each view:
  - Its *view name* and *definition.*

In addition, statistics about tables and indexes are stored in the system catalogs and updated periodically *(not* every time the underlying tables are modified). The following information is commonly stored:

- Cardinality: The number of tuples *NTuples(R)* for each table *R.*
- Size: The number of pages *NPages(R)* for each table *R.*
- Index Cardinality: The number of distinct key values *NKeys(I)* for each index *I.*
- Index Size: The nUluber of pages *INPages(I)* for each index *I.* (For a B+ tree index *I,* we take *INPagcs* to be the number of leaf pages.)

- **Index Height:** The number of nonleaf levels *IHe'ight(I)* for each tree index *I*.
- **Index Range:** The minimum present key value *ILow(I)* and the maximum present key value *IHigh(I)* for each index *I*.

We assume that the database architecture presented in Chapter 1 is used. Further, we assume that each file of records is implemented as a separate file of pages. Other file organizations are possible, of course. For example, a page file can contain pages that store records from more than one record file. If such a file organization is used, additional statistics must be maintained, such as the fraction of pages in a file that contain records from a given collection of records.

The catalogs also contain information about *users*, such as accounting information and *authorization* information (e.g., Joe User can modify the Reserves table but only read the Sailors table).

How Catalogs are Stored

An elegant aspect of a relational DBMS is that the system catalog is itself a collection of tables. For example, we might store information about the attributes of tables in a catalog table called Attribute_Cat:

Attribute_Cat(*attr_name:* string, *rel_name:* string,
type: string, *position:* integer)

Suppose that the database contains the two tables that we introduced at the begining of this chapter:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)
Reserves(sid: integer, *bid:* integer, *day:* dates, *mame:* string)

Figure 12.1 shows the tuples in the Attribute_Cat table that describe the attributes of these two tables. Note that in addition to the tuples describing Sailors and Reserves, other tuples (the first four listed) describe the four attributes of the Attribute_Cat table itself! These other tuples illustrate an important Point: the catalog tables describe all the tables in the database, *including* the catalog tables themselves. When information about a table is needed, it is obtained from the system catalog. Of course, at the implementation level, whenever the DBMS needs to find the schema of a *catalog* table, the code that retrieves this information must be handled specially. (Otherwise, the code has to retrieve this information from the catalog tables without, presumably, knowing the schema of the catalog tables.)

attr_name	*rel*	*type*	
attr_name	Attribute_Cat	string	1
reLname	Attribute_Cat	string	2
type	Attribute_Cat	string	3
position	Attribute_Cat	integer	4
sid	Sailors	integer	1
sname	Sailors	string	2
rating	Sailors	integer	3
age	Sailors	real	4
sid	Reserves	integer	1
bid	Reserves	integer	2
day	Reserves	dates	3
rname	Reserves	string	4

Figure 12.1 An Instance of the Attribute_Cat Relation

The fact that the system catalog is also a collection of tables is very useful. For example, catalog tables can be queried just like any other table, using the query language of the DBMS! Further, all the techniques available for implementing and managing tables apply directly to catalog tables. The choice of catalog tables and their schemas is not unique and is made by the implementor of the DBMS. Real systems vary in their catalog schema design, but the catalog is always implemented as a collection of tables, and it essentially describes all the data stored in the database.[1]

12.2 INTRODUCTION TO OPERATOR EVALUATION

Several alternative algorithms are available for implementing each relational operator, and for most operators no algorithm is universally superior. Several factors influence which algorithm performs best, including the sizes of the tables involved, existing indexes and sort orders, the size of the available buffer pool, and the buffer replacement policy.

In this section, we describe some common techniques used in developing evaluation algorithms for relational operators, and introduce the concept of *access paths*, which are the different ways in which rows of a table can be retrieved.

[1]Some systems may store additional information in a non-relational form. For example, a system with a sophisticated query optimizer may maintain histograms or other statistical information about the distribution of values in certain attributes of a table. We can think of such information, when it is maintained, as a supplement to the catalog tables.

12.2.1 Three Common Techniques

The algorithms for various relational operators actually have a lot in common. A few simple techniques are used to develop algorithms for each operator:

- Indexing: If a selection or join condition is specified, use an index to examine just the tuples that satisfy the condition.
- Iteration: Examine all tuples in an input table, one after the other. If we need only a few fields from each tuple and there is an index whose key contains all these fields, instead of examining data tuples, we can scan all index data entries. (Scanning all data entries sequentially makes no use of the index's hash- or tree-based search structure; in a tree index, for example, we would simply examine all leaf pages in sequence.)
- Partitioning: By partitioning tuples on a sort key, we can often decompose an operation into a less expensive collection of operations on partitions. *Sorting* and *hashing* are two commonly used partitioning techniques.

We discuss the role of indexing in Section 12.2.2. The iteration and partitioning techniques are seen in Section 12.3.

12.2.2 Access Paths

An access path is a way of retrieving tuples from a table and consists of either (1) a file scan or (2) an index plus a *matching* selection condition. Every relational operator accepts one or more tables as input, and the access methods used to retrieve tuples contribute significantly to the cost of the operator.

Consider a simple selection that is a conjunction of conditions of the form $attr$ op $value$, where op is one of the comparison operators $<$, $\leq$, $=$, $\neq$, $\geq$, or $>$. Such selections are said to be in conjunctive normal form (CNF), and each condition is called a conjunct.[2] Intuitively, an index matches a selection condition if the index can be used to retrieve just the tuples that satisfy the condition.

- A hash index matches a CNF selection if there is a term of the form $attribute$=$value$ in the selection for each attribute in the index's search key.
- A tree index matches a CNF selection if there is a term of the form $attribute$ op $value$ for each attribute in a *prefix* of the index's search key. ($\langle a \rangle$ and $\langle a, b \rangle$ are prefixes of key (a,b,e), but $\langle a, c \rangle$ and $\langle b, c \rangle$ are not.)

[2] We consider more complex selection conditions in Section 14.2.

Note that **op** can be any comparison; it is not restricted to he equality as it is for matching selections on a hash index.

An index can match some subset of the conjuncts in a selection condition (in CNP), even though it does not match the entire condition. We refer to the conjuncts that the index matches as the **primary conjuncts** in the selection.

The following examples illustrate access paths.

- If we have a hash index H on the search key $\langle rname,bid,sid\rangle$, we can use the index to retrieve just the Sailors tuples that satisfy the condition *rnarne='Joe'l\ bid=5 1\ sid=3.* The index matches the entire condition *rname='Joe' 1\ bid=5* $\wedge$ *sid=* 3. On the other hand, if the selection condition is *rname='Joe' 1\ bid=5,* or some condition on *date,* this index does not match. That is, it cannot be used to retrieve just the tuples that satisfy these conditions.

  In contrast, if the index were a B+ tree, it would match both *rname='Joe' 1\ bid=51\ 8id=3* and *mame='Joe' 1\ bid=5.* However, it would not match *bid=5 1\ sid=3* (since tuples are sorted primarily by *rnarne).*

- If we have an index (hash or tree) on the search key $\langle bid,sid\rangle$ and the selection condition *rname='Joe' 1\ bid=5 1\ sid=3,* we can use the index to retrieve tuples that satisfy *bid=51\ sid=3;* these are the primary conjuncts. The fraction of tuples that satisfy these conjuncts (and whether the index is clustered) determines the number of pages that are retrieved. The additional condition on *Tna7ne* must then be applied to each retrieved tuple and will eliminate some of the retrieved tuples from the result.

- If we have an index on the search key $\langle bid, sid\rangle$ and we also have a B+ tree index on *day,* the selection condition *day* < *8/9/2002 1\ bid=5* $\wedge$ *sid=3* offers us a choice. Both indexes match (part of) the selection condition, and we can use either to retrieve Reserves tuples. \Vhichever index we use, the conjuncts in the selection condition that are not matched by the index (e.g., *bid=51\ sid=3* if we use the B+ tree index on *day*) must be checked for each retrieved tuple.

Selectivity of Access Paths

The selectivity of an access path is the number of pages retrieved (index pages plus data pages) if we usc this access path to retrieve all desired tuples. If a table contains an index that matches a given selection, there are at least two access paths: the index and a scan of the data file. Sometimes, of course, we can scan the index itself (rather than scanning the data file or using the index to probe the file), giving us a third access path.

The most selective access path is the one that retrieves the fewest pages; using the most selective access path minimizes the cost of data retrieval. The selectivity of an access path depends on the primary conjuncts in the selection condition (with respect to the index involved). Each conjunct acts as a filter on the table. The fraction of tuples in the table that satisfy a given conjunct is called the reduction factor. 'When there are several primary conjuncts, the fraction of tuples that satisfy all of them can be approximated by the product of their reduction factors; this effectively treats them as independent filters, and while they may not actually be independent, the approximation is widely used in practice.

Supose we have a hash index H on Sailors with search key $\langle rname, bid, sid \rangle$, and we are given the selection condition *rname='Joe' 1\ bid=5 1\ sid=3.* We can use the index to retrieve tuples that satisfy all three conjuncts. The catalog contains the number of distinct key values, $NKeys(H)$, in the hash index, as well as the number of pages, $NPages$, in the Sailors table. The fraction of pages satisfying the primary conjuncts is $Npages(Sailors) \cdot \frac{1}{NKeys(H)}$.

If the index has search key $\langle bid, sid \rangle$, the primary conjuncts are *bid=51\ sid=3.* If we know the number of distinct values in the *bid* column, we can estimate the reduction factor for the first conjunct. This information is available in the catalog if there is an index with *bid* as the search key; if not, optimizers typically use a default value such as 1/10. Multiplying the reduction factors for *bid=5* and *sid=3* gives us (under the simplifying independence assumption) the fraction of tuples retrieved; if the index is clustered, this is also the fraction of pages retrieved. If the index is not clustered, each retrieved tuple could be on a different page. (Review Section 8.4 at this time.)

We estimate the reduction factor for a range condition such as *day> 8/9/2002* by assuming that values in the column are uniformly distributed. If there is a B+ tree T with key day, the reduction factor is $\frac{High(T) - value}{High(T) - Low(T)}$.

12.3 ALGORITHMS FOR RELATIONAL OPERATIONS

We now briefly discuss evaluation algorithms for the main relational operators. While the important ideas are introduced here, a more in-depth treatment is deferred to Chapter 14. As in Chapter 8, we consider only I/O costs and measure I/O costs in terms of the number of page I/Os. In this chapter, we use detailed examples to illustrate how to compute the cost of an algorithm. Although we do not present rigorous cost formulas in this chapter, the reader should be able to apply the underlying icleas to do cost calculations on other similar examples.

12.3.1 Selection

The selection operation is a simple retrieval of tuples from a table, and its implementation is essentially covered in our discussion of access paths. To summarize, given a selection of the form $\sigma_{R.attr\ \mathbf{op}\ value}(R)$, if there is no index on *R.attr,* we have to scan *R.*

If one or more indexes on *R* match the selection, we can use the index to retrieve matching tuples, and apply any remaining selection conditions to further restrict the result set. As an example, consider a selection of the form *rname* < *'C%'* on the Reserves table. Assuming that names are uniformly distributed with respect to the initial letter, for simplicity, we estimate that roughly 10% of Reserves tuples are in the result. This is a total of 10,000 tuples, or 100 pages. If we have a clustered B+ tree index on the *rname* field of Reserves, we can retrieve the qualifying tuples with 100 l/Os (plus a few l/Os to traverse from the root to the appropriate leaf page to start the scan). However, if the index is unclustered, we could have up to 10,000 l/Os in the worst case, since each tuple could cause us to read a page.

As a rule of thumb, it is probably cheaper to simply scan the entire table (instead of using an unclustered index) if over 5% of the tuples are to be retrieved.

Sec Section 14.1 for more details on implementation of selections.

12.3.2 Projection

The projection operation requires us to drop certain fields of the input, which is easy to do. The expensive aspect of the operation is to ensure that no duplicates appear in the result. For example, if we only want the *sid* and *bid* fields from Reserves, we could have duplicates if a sailor has reserved a given boat on several days.

If duplicates need not be eliminated (e.g., the DISTINCT keyword is not included in the SELECT clause), projection consists of simply retrieving a subset of fields from each tuple of the input table. This can be accomplished by simple iteration on either the table or an index whose key contains all necessary fields. (Note that we do not care whether the index is clustered, since the values we want are in the data entries of the index itself!)

If we have to eliminate duplicates, we typically have to use partitioning. Suppose we want to obtain $\langle sid, bid \rangle$ by projecting from Reserves. We can partition by (1) scanning H.eserves to obtain *(sid, bid)* pairs and (2) sorting these pairs

using ⟨*sid, bid*⟩ as the sort key. We can then scan the sorted pairs and easily discard duplicates, which are now adjacent.

Sorting large disk-resident datasets is a very important operation in database systems, and is discussed in Chapter 13. Sorting a table typically requires two or three passes, each of which reads and writes the entire table.

The projection operation can be optimized by combining the initial scan of Reserves with the scan in the first pass of sorting. Similarly, the scanning of sorted pairs can be combined with the last pass of sorting. With such an optimized implemention, projection with duplicate elimination requires (1) a first pass in which the entire table is scanned, and only pairs ⟨*sid, bid*⟩ are written out, and (2) a final pass in which all pairs are scanned, but only one copy of each pair is written out. In addition, there might be an intermediate pass in which all pairs are read from and written to disk.

The availability of appropriate indexes can lead to less expensive plans than sorting for duplicate elimination. If we have an index whose search key contains all the fields retained by the projection, we can sort the data entries in the index, rather than the data records themselves. If all the retained attributes appear in a prefix of the search key for a clustered index, we can do even better; we can simply retrieve data entries using the index, and duplicates are easily detected since they are adjacent. These plans are further examples of *index-only* evaluation strategies, which we discussed in Section 8.5.2.

See Section 14.3 for more details on implementation of projections.

12.3.3 Join

Joins are expensive operations and very common. Therefore, they have been widely studied, and systems typically support several algorithms to carry out joins.

Consider the join of Reserves and Sailors, with the join conclition *Reserves.sid* = *Sailors.sid*. Suppose that one of the tables, say Sailors, has an index on the *sid* column. We can scan Reserves and, for each tuple, use the index to *pTObe* Sailors for matching tuples. This approach is called index nested loops join.

Suppose that we have a hash-based index using Alternative (2) on the *sid* attribute of Sailors and that it takes about 1.2 1/0s on average[3] to retrieve the appropriate page of the index. Since *sid* is a key for Sailors, we have at

:IThis is a typical cost for hash-based indexes.

most one matching tuple, Indeed, *sid* in Reserves is a foreign key referring to Sailors, and therefore we have *exactly* one matching Sailors tuple for each Reserves tuple, Let us consider the cost of scanning Reserves and using the index to retrieve the matching Sailors tuple for each Reserves tuple, The cost of scanning Reserves is 1000. There are 100 * 1000 tuples in Reserves. For each of these tuples, retrieving the index page containing the rid of the matching Sailors tuple costs 1.2 I/Os (on average); in addition, we have to retrieve the Sailors page containing the qualifying tuple, Therefore, we have 100,000 * (1 + 1.2) I/Os to retrieve matching Sailors tuples. The total cost is 221,000 I/Os.[4]

If we do not have an index that matches the join condition on either table, we cannot use index nested loops, In this case, we can sort both tables on the join column, and then scan them to find matches. This is called sort-merge join.. Assuming that we can sort Reserves in two passes, and Sailors in two passes as well, let us consider the cost of sort-merge join. Consider the join of the tables Reserves and Sailors. Because we read and write Reserves in each pass, the sorting cost is 2 * 2 * 1000 = 4000 I/Os. Similarly, we can sort Sailors at a cost of 2 * 2 * 500 = 2000 I/Os. In addition, the second phase of the sort-merge join algorithm requires an additional scan of both tables. Thus the total cost is 4000 + 2000 + 1000 + 500 = 7500 I/Os.

Observe that the cost of sort-merge join, which does not require a pre-existing index, is lower than the cost of index nested loops join, In addition, the result of the sort-merge join is sorted on the join column(s). Other join algorithms that do not rely on an existing index and are often cheaper than index nested loops join are also known *(block nested loops* and *hash* joins; see Chapter 14). Given this, why consider index nested loops at all?

Index nested loops has the nice property that it is incremental. The cost of our example join is incremental in the number of Reserves tuples that we process. Therefore, if some additional selection in the query allows us to consider only a small subset of Reserves tuples, we can avoid computing the join of Reserves and Sailors in its entirety. For instance, suppose that we only want the result of the join for boat 101, and there are very few such reservations. For each such Reserves tuple, we probe Sailors, and we are clone. If we use sort-merge join, on the other hand, we have to scan the entire Sailors table at least once, and the cost of this step alone is likely to be much higher than the entire cost of index nested loops join.

Observe that the choice of index nested loops join is based on considering the query as a whole, including the extra selection all Reserves, rather than just

[4]As an exercise, the reader should write formulas for the cost estimates in this example in terms of the properties e.g., NPages-of the tables and indexes involved.

the join operation by itself. This leads us to our next topic, query optimization, which is the process of finding a good plan for an entire query.

See Section 14.4 for more details.

12.3.4 Other Operations

A SQL query contains group-by and aggregation in addition to the basic relational operations. Different query blocks can be combined with union, set-difference, and set-intersection.

The expensive aspect of set operations such as union and intersection is duplicate elimination, just like for projection. The approach used to implement projection is easily adapted for these operations as well. See Section 14.5 for more details.

Group-by is typically implemented through sorting. Sometimes, the input table has a tree index with a search key that matches the grouping attributes. In this case, we can retrieve tuples using the index in the appropriate order without an explicit sorting step. Aggregate operations are carried out using temporary counters in main memory as tuples are retrieved. See Section 14.6 for more details.

12.4 INTRODUCTION TO QUERY OPTIMIZATION

Query optimization is one of the most important tasks of a relational DBMS. One of the strengths of relational query languages is the wide variety of ways in which a user can express and thus the system can evaluate a query. Although this flexibility makes it easy to write queries, good performance relies greatly on the quality of the query optimizer—a given query can be evaluated in many ways, and the difference in cost between the best and worst plans may be several orders of magnitude. Realistically, we cannot exped to always find the best plan, but we expect to consistently find a plan that is quite good.

A more detailed view of the query optimization and execution layer in the DBMS architecture from Section 1.8 is shown in Figure 12.2. Queries are parsed and then presented to a **query optimizer,** which is responsible for identifying an efficient execution plan. The optimizer generates alternative plans and chooses the plan wit.h the least estimated cost.

The space of plans considered by a typical relational query optimizer can be understood by recognizing that *a query is essentially treated as a* σ - π - $\bowtie$ *algebra expression,* with the remaining operations (if any, in a given query)

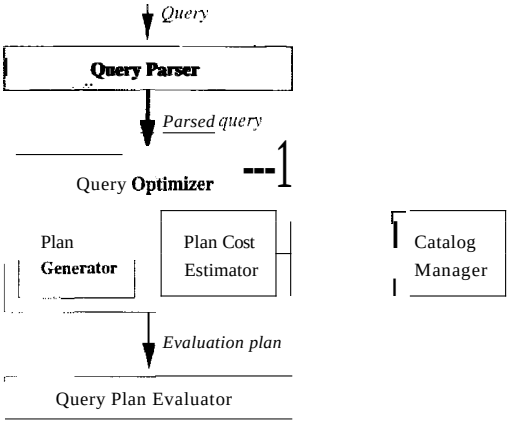


Figure 12.2 Query Parsing, Optimization, and Execution

CommercialOptimizers: Current relational DBMS optimizers are very complex pieces of software with many closely guarded details, and they typically represent 40 to 50 man-years of development effort!

carried out on the result of the σ - π— [x] expression. Optimizing such a relational algebra expression involves two basic steps:

- Enumerating alternative plans for evaluating the expression. Typically, an optimizer considers a subset of all possible plans because the number of possible plans is very large.
- Estimating the cost of each enumerated plan and choosing the plan with the lowest estimated cost.

In this section we lay the foundation for our discussion of query optimization by introducing evaluation plans.

12.4.1 Query Evaluation Plans

A query evaluation plan (or simply plan) consists of an extended relational algebra tree, with additional annotations at each node indicating the access methods to use for each table and the implementation method to use for each relational operator.

Consider the following SQL query:

```
SELECT S.sname
FROM   Reserves R, Sailors S
WHERE  R.sid = S.sid
       AND R.bid = 100 AND S.rating > 5
```

This query can be expressed in relational algebra as follows:

$$\pi_{sname}(\sigma_{bid=100 \wedge rating>5}(Reserves \bowtie_{sid=sid} Sailors))$$

This expression is shown in the form of a tree in Figure 12.3. The algebra expression partially specifies how to evaluate the query-owe first compute the natural join of Reserves and Sailors, then perform the selections, and finally project the *snarne* field.

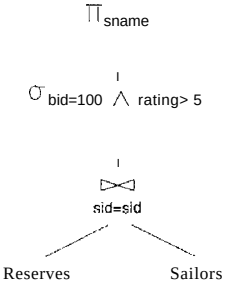


Figure 12.3 Query Expressed as a Relational Algebra Tree

To obtain a fully specified evaluation plan, we must decide on an implementation for each of the algebra operations involved. For example, we can use a page-oriented simple nested loops join with Reserves as the outer table and apply selections and projections to each tuple in the result of the join as it is produced; the result of the join before the selections and projections is never stored in its entirety. This query evaluation plan is shown in Figure 12.4.

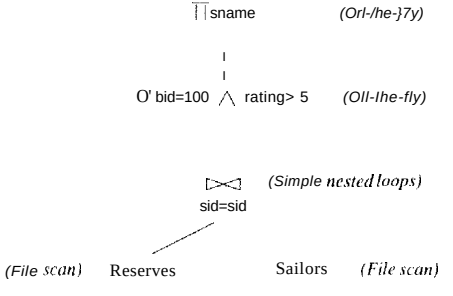


Figure 12.4 Query Evaluation Plan for Sample Query

In drawing the query evaluation plan, we have used the convention that the *outer table* is the *left child* of the join operator. We adopt this convention henceforth.

12.4.2 Multi-operator Queries: Pipelined Evaluation

When a query is composed of several operators, the result of one operator is sometimes pipelined to another operator without creating a temporary table to hold the intermediate result. The plan in Figure 12.4 pipelines the output of the join of Sailors and Reserves into the selections and projections that follow. Pipelining the output of an operator into the next operator saves the cost of writing out the intermediate result and reading it back in, and the cost savings can be significant. If the output of an operator is saved in a temporary table for processing by the next operator, we say that the tuples are materialized. Pipelined evaluation has lower overhead costs than materialization and is chosen whenever the algorithm for the operator evaluation permits it.

There are many opportunities for pipelining in typical query plans, even simple plans that involve only selections.. Consider a selection query in which only part of the selection condition matches an index. We can think of such a query as containing *two* instances of the selection operator: The first contains the primary, or matching, part of the original selection condition, and the second contains the rest of the selection condition. We can evaluate such a query by applying the primary selection and writing the result to a temporary table and then applying the second selection to the temporary table. In contrast, a pipelined evaluation consists of applying the second selection to each tuple in the result of the primary selection as it is produced and adding tuples that qualify to the final result. When the input table to a unary operator (e.g., selection or projection) is pipelined into it, we sometimes say that the operator is applied on-the-fly.

As a second and more general example, consider a join of the form $(A \bowtie B) \bowtie C$, shown in Figure 12.5 as a tree of join operations.

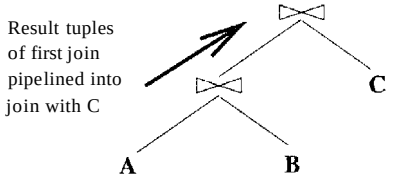


Figure 12.5 A Query Tree Illustrating Pipelilling

Both joins can be evaluated in pipelined fashion using some version of a nested loops join. Conceptually, the evaluation is initiated from the root, and the node joining A and B produces tuples as and when they are requested by its parent node. 'When the root node gets a page of tuples from its left child (the outer table), all the matching inner tuples are retrieved (using either an index or a scan) and joined with matching outer tuples; the current page of outer tuples is then discarded. A request is then made to the left child for the next page of tuples, and the process is repeated. Pipelined evaluation is thus a *control strategy* governing the rate at which different joins in the plan proceed. It has the great virtue of not writing the result of intermediate joins to a temporary file because the results are produced, consumed, and discarded one page at a time.

12.4.3 The Iterator Interface

A query evaluation plan is a tree of relational operators and is executed by calling the operators in some (possibly interleaved) order. Each operator has one or more inputs and an output, which are also nodes in the plan, and tuples must be passed between operators according to the plan's tree structure.

To simplify the code responsible for coordinating the execution of a plan, the relational operators that form the nodes of a plan tree (which is to be evaluated using pipelining) typically support a uniform **iterator** interface, hiding the internal implementation details of each operator. The iterator interface for an operator includes the functions open, geLnext, and close. The *open* function initializes the state of the iterator by allocating buffers for its inputs and output, and is also used to pass in arguments such as selection conditions that modify the behavior of the operator. The code for the *get_next* function calls the *get_next* function on each input node and calls operator-specific code to process the input tuples. The output tuples generated by the processing are placed in the output buffer of the operator, and the state of the iterator is updated to keep track of how much input has been consumed. When all output tuples have been produced through repeated calls to *get_next*, the *close* function is called (by the code that initiated execution of this operator) to deallocate state information.

The iterator interface supports pipelining of results naturally: the decision to pipeline or materialize input tuples is encapsulated in the operator-specific code that processes input tuples. If the algorithm implemented for the operator allows input tuples to be processed completely when they are received, input tuples are not Inaterialized and the evaluation is pipelined. If the algorithm examines the same input tuples several times, they are materialized. This

decision, like other details of the operator's implementation, is hidden by the iterator interface for the operator.

The iterator interface is also used to encapsulate access methods such as B+ trees and hash-based indexes. Externally, access methods can be viewed simply as operators that produce a stream of output tuples. In this case, the *open* function can be used to pass the selection conditions that match the access path.

12.5 ALTERNATIVE PLANS: A MOTIVATING EXAMPLE

Consider the example query from Section 12.4. Let us consider the cost of evaluating the plan shown in Figure 12.4. We ignore the cost of writing out the final result since this is common to all algorithms, and does not affect their relative costs. The cost of the join is 1000 + 1000 * 500 = 501,000 page I/Os. The selections and the projection are done on-the-fly and do not incur additional I/Os. The total cost of this plan is therefore 501,000 page I/Os. This plan is admittedly naive; however, it is possible to be even more naive by treating the join as a cross-product followed by a selection.

We now consider several alternative plans for evaluating this query. Each alternative improves on the original plan in a different way and introduces some optimization ideas that are examined in more detail in the rest of this chapter.

12.5.1 Pushing Selections

A join is a relatively expensive operation, and a good heuristic is to reduce the sizes of the tables to be joined as much as possible. One approach is to apply selections early; if a selection operator appears after a join operator, it is worth examining whether the selection can be 'pushed' ahead of the join. As an example, the selection *bid=100* involves only the attributes of Reserves and can be applied to Reserves *before* the join. Similarly, the selection *rating> 5* involves only attributes of Sailors and can be applied to Sailors before the join. Let us suppose that the selections are performed using a simple file scan, that the result of each selection is written to a temporary table on disk, and that the temporary tables are then joined using a sort-merge join. The resulting query evaluation plan is shown in Figure 12.6.

Let us assume that five buffer pages are available and estimate the cost of this query evaluation plan. (It is likely that more buffer pages are available in practice. We chose a small number simply for illustration in this example.) The cost of applying *bid=100* to Reserves is the cost of scanning Reserves (1000 pages) plus the cost of writing the result to a temporary table, say T1.

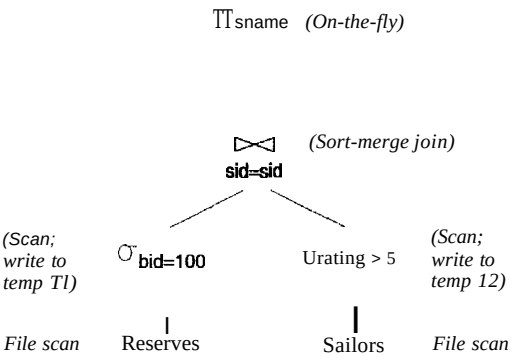


Figure 12.6 A Second Query Evaluation Plan

(Note that the cost of writing the temporary table cannot be ignored-we can ignore only the cost of writing out the *final* result of the query, which is the only component of the cost that is the same for all plans.) To estimate the size of T1, we require additional information. For example, if we assume that the maximum number of reservations of a given boat is one, just one tuple appears in the result. Alternatively, if we know that there are 100 boats, we can assume that reservations are spread out uniformly across all boats and estimate the number of pages in T1 to be 10. For concreteness, assume that the number of pages in T1 is indeed 10.

The cost of applying *rating* > 5 to Sailors is the cost of scanning Sailors (500 pages) plus the cost of writing out the result to a temporary table, say, T2. If we assume that ratings are uniformly distributed over the range 1 to 10, we can approximately estimate the size of T2 as 250 pages.

To do a sort-merge join of T1 and T2, let us assume that a straightforward implementation is used in which the two tables are first completely sorted and then merged. Since five buffer pages are available, we can sort T1 (which has 10 pages) in two passes. Two runs of five pages each are produced in the first pass and these are merged in the second pass. In each pass, we read and write 10 pages; thus, the cost of sorting T1 is $2 * 2 * 10 = 40$ page I/Os. We need four passes to sort T2, which has 250 pages. The cost is $2 * 4 * 250 = 2000$ page I/Os. To merge the sorted versions of T1 and T2, we need to scan these tables, and the cost of this step is $10 + 250 = 260$. The final projection is done on-the-fly, and by convention we ignore the cost of writing the final result.

The total cost of the plan shown in Figure 12.6 is the sum of the cost of the selection (1000+10+500+250 = 1760) and the cost of the join (40+2000+260 = 23(0), that is, 4060 page I/Os.

Sort-merge join is one of several join methods. We may be able to reduce the cost of this plan by choosing a different join method. As an alternative, suppose that we used block nested loops join instead of sort-merge join. Using T1 as the outer table, for every three-page block of T1, we scan all of T2; thus, we scan T2 four times. The cost of the join is therefore the cost of scanning T1 (10) plus the cost of scanning T2 (4 * 250 = 1000). The cost of the plan is now 1760 + 1010 = 2770 page I/Os.

A further refinement is to push the projection, just like we pushed the selections past the join. Observe that only the *sid* attribute of T1 and the *sid* and *sname* attributes of T2 are really required. As we scan Reserves and Sailors to do the selections, we could also eliminate unwanted columns. This on-the-fly projection reduces the sizes of the temporary tables T1 and T2. The reduction in the size of T1 is substantial because only an integer field is retained. In fact, T1 now fits within three buffer pages, and we can perform a block nested loops join with a single scan of T2. The cost of the join step drops to under 250 page I/Os, and the total cost of the plan drops to about 2000 I/Os.

12.5.2 Using Indexes

If indexes are available on the Reserves and Sailors tables, even better query evaluation plans may be available. For example, suppose that we have a clustered static hash index on the *bid* field of Reserves and another hash index on the *sid* field of Sailors. We can then use the query evaluation plan shown in Figure 12.7.

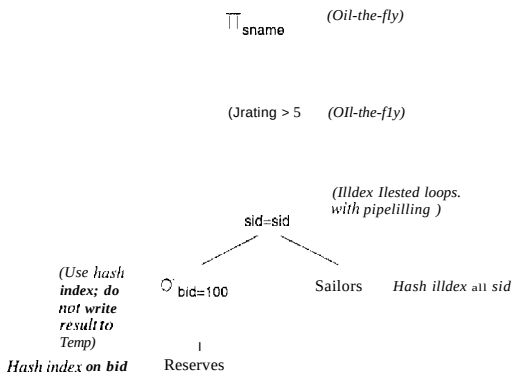


Figure 12.7 A Query Evaluation Plan Using Indexes

The selection *bid*=100 is performed on Reserves by using the hash index on *bid* to retrieve only matching tuples. As before, if we know that 100 boats are available and assume that reservations are spread out uniformly across all boats,

we can estimate the number of selected tuples to be 100, 000/100 = l000. Since the index on *b'id* is clustered, these 1000 tuples appear consecutively within the same bucket; therefore, the cost is 10 page l/Os.

:For each selected tuple, we retrieve matching Sailors tuples using the hash index on the *sid* field; selected Reserves tuples are not materialized and the join is pipelined. For each tuple in the result of the join, we perform the selection *rating*>5 and the projection of *sname* on-the-fly. There are several important points to note here:

1. Since the result of the selection on Reserves is not materialized, the optimization of projecting out fields that are not needed subsequently is unnecessary (and is not used in the plan shown in Figure 12.7).
2. The join field *sid* is a key for Sailors. Therefore, at most one Sailors tuple matches a given Reserves tuple. The cost of retrieving this matching tuple depends on whether the directory of the hash index on the *sid* column of Sailors fits in memory and on the presence of overflow pages (if any). However, the cost does *not* depend on whether this index is clustered because there is at most one matching Sailors tuple and requests for Sailors tuples are made in random order by *sid* (because Reserves tuples are retrieved by *bid* and are therefore considered in random order by *sid*). For a hash index, 1.2 page l/Os (on average) is a good estimate of the cost for retrieving a data entry. Assuming that the *sid* hash index on Sailors uses Alternative (1) for data entries, 1.2 l/Os is the cost to retrieve a matching Sailors tuple (and if one of the other two alternatives is used, the cost would be 2,2 l/Os).
3. We have chosen not to push the selection *rating*>5 ahead of the join, and there is an important reason for this decision. If we performed the selection before the join, the selection would involve scanning Sailors, assuming that no index is available on the *rating* field of Sailors. Further, whether or not such an index is available, once we apply such a selection, we have no index on the *sid* field of the result of the selection (unless we choose to build such an index solely for the sake of the subsequent join). Thus, pushing selections ahead of joins is a good heuristic, but not always the best strategy. Typically, as in this example, the existence of useful indexes is the reason a selection is *not* pushed. (Otherwise, selections are pushed.)

Let us estimate the cost of the plan shown in Figure 12.7. The selection of Reserves tuples costs 10 l/Os, as we saw earlier. There are 1000 such tuples, and for each, the cost of finding the matching Sailors tuple is 1.2 l/Os, on average. The cost of this step (the join) is therefore 1200 l/Os. All remaining selections and projections are performed on-the-fly. The total cost of the plan is 1210 l/Os.

As noted earlier, this plan does not utilize clustering of the Sailors index. The plan can be further refined if the index on the *sid* field of Sailors is clustered. Suppose we materialize the result of performing the selection *bid*=100 on Reserves and sort this temporary table. This table contains 10 pages. Selecting the tuples costs 10 page I/Os (as before), writing out the result to a temporary table costs another 10 I/Os, and with five buffer pages, sorting this temporary costs 2 * 2 * 10 = 40 I/Os. (The cost of this step is reduced if we push the projection on *sid*. The *sid* column of materialized Reserves tuples requires only three pages and can be sorted in memory with five buffer pages.) The selected Reserves tuples can now be retrieved in order by *sid*.

If a sailor has reserved the same boat many times, all corresponding Reserves tuples are now retrieved consecutively; the matching Sailors tuple will be found in the buffer pool on all but the first request for it. This improved plan also demonstrates that pipelining is not always the best strategy.

The combination of pushing selections and using indexes illustrated by this plan is very powerful. If the selected tuples from the outer table join with a single inner tuple, the join operation may become trivial, and the performance gains with respect to the naive plan in Figure 12.6 are even more dramatic. The following variant of our example query illustrates this situation:

```
SELECT  S.sname
FROM    Reserves R, Sailors S
WHERE   Rsid = S.sid
        AND R.bid = 100 AND S.rating > 5
        AND Rday = '8/9/2002'
```

A slight variant of the plan shown in Figure 12.7, designed to answer this query, is shown in Figure 12.8. The selection *day='8/9/2002'* is applied on-the-fly to the result of the selection *bid=100* on the Reserves table.

Suppose that *bid* and *day* form a key for Reserves. (Note that this assumption differs from the schema presented earlier in this chapter.) Let us estimate the cost of the plan shown in Figure 12.8. The selection *bid=100* costs 10 page I/Os, as before, and the additional selection *day='8/9/2002'* is applied on-the-fly, eliminating all but (at most) one Reserves tuple. There is at most one rnatching Sailors tuple, and this is retrieved in 1.2 I/Os (an average value). The selection on *rating* and the projection on *sname* are then applied on-the-fly at no additional cost. The total cost of the plan in Figure 12.8 is thus about 11 I/Os. In contrast, if we modify the naive plan in Figure 12.6 to perform the additional selection on *day* together with the selection *bid=100*, the cost remains at 501,000 I/Os.

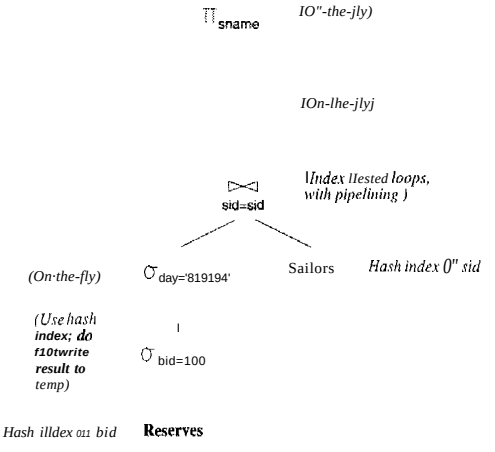


Figure 12.8 A Query Evaluation Plan for the Second Example

12.6 WHAT A TYPICAL OPTIMIZER DOES

A relational query optimizer uses relational algebra equivalences to identify many equivalent expressions for a given query. For each such equivalent version of the query, all available implementation techniques are considered for the relational operators involved, thereby generating several alternative queryevaluation plans. The optimizer estimates the cost of each such plan and chooses the one with the lowest estimated cost.

12.6.1 Alternative Plans Considered

Two relational algebra expressions over the same set of input tables are said to be **equivalent** if they produce the same result on all instances of the input tables. Relational algebra equivalences playa central role in identifying alternative plans.

Consider a basic SQL query consisting of a SELECT clause, a FROM clause, and a WHERE clause, This is easily represented as an algebra expression; the fields mentioned in the SELECT are projected from the cross-product of tables in the FROM clause, after applying the selections in the WHERE clause. The use of equivalences enable us to convert this initial representation into equivalent expressions. In particular:

- Selections and cross-products can be combined into joins.
- Joins can be extensively reordered.

- Selections and projections, which reduce the size of the input, can be "pushed" ahead of joins.

The query discussed in Section 12.5 illustrates these points; pushing the selection in that query ahead of the join yielded a dramatically better evaluation plan. \Ve discuss relational algebra equivalences in detail in Section 15.3.

Left-Deep Plans

Consider a query of the form $A \bowtie B \bowtie C \bowtie D$; that is, the natural join of four tables. Three relational algebra operator trees that are equivalent to this query (based on algebra equivalences) are shown in Figure 12.9. By convention, the left child of a join node is the outer table and the right child is the inner table. By adding details such as the join method for each join node, it is straightforward to obtain several query evaluation plans from these trees.

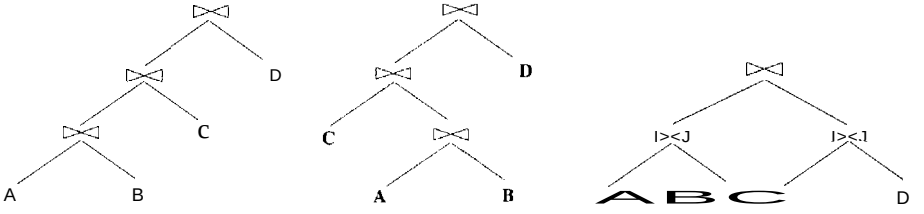


Figure 12.9 Three Join Trees

The first two trees in Figure 12.9 are examples of linear trees. In a linear tree, at least one child of a join node is a base table. The first tree is an example of a left-deep tree-the *right* child of each join node is a base table. The third tree is an example of a non-linear or bushy tree.

Optimizers typically use a dynamic-programming approach (see Section 15.4.2) to efficiently search the class of aU left-deep plans. The second and third kinds of trees are therefore never considered. Intuitively, the first tree represents a plan in which we join A and B first, then join the result with C, then join the result with D. There are 23^5 other left-deep plans that differ only in the order that tables are joined. If any of these plans has selection and projection conditions other than the joins themselves, these conditions are applied as early as possible (consitent with algebra equivalences) given the choice of a join order for the tables.

Of course, this decision rules out many alternative plans that may cost less than the best plan using a left-deep tree; we have to live with the fact that

[5]The reader should think through the number 23 in this example.

the optimizer will never find such plans. There are two main reasons for this decision to concentrate on **left-deep plans,** or plans based on left-deep trees:

1. As the number of joins increases, the number of alternative plans increases rapidly and it becomes necessary to prune the space of alternative plans.
2. Left-deep trees allow us to generate all **fully pipelined** plans; that is, plans in which all joins are evaluated using pipelining. (Inner tables must always be materialized because we must examine the entire inner table for each tuple of the outer table. So, a plan in which an inner table is the result of a join forces us to materialize the result of that join.)

12.6.2 Estimating the Cost of a Plan

The cost of a plan is the sum of costs for the operators it contains. The cost of individual relational operators in the plan is estimated using information, obtained from the system catalog, about properties (e.g., size, sort order) of their input tables. We illustrated how to estimate the cost of single-operator plans in Sections 12.2 and 12.3, and how to estimate the cost of multi-operator plans in Section 12.5.

If we focus on the metric of I/O costs, the cost of a plan can be broken down into three parts: (1) reading the input tables (possibly rnultiple times in the case of some join and sorting algorithms), (2) writing intermediate tables, and (possibly) (3) sorting the final result (if the query specifies duplicate elimination or an output order). The third part is common to all plans (unless one of the plans happens to produce output in the required order), and, in the common case that a fully-pipelined plan is chosen, no intermediate tables are written.

Thus, the cost for a fully-pipelined plan is dominated by part (1). This cost depends greatly on the access paths used to read input tables; of course, access paths that are used repeatedly to retrieve matching tuples in a join algorithm are especially important.

For plans that are not fully pipelined, the cost of rnaterializing temporary tables can be significant. The cost of materializing an intermediate result depends on its size, and the size also infiuences the cost of the operator for which the temporary is an input table. The number of tuples in the result of a selection is estimated by multiplying the input size by the reduction factor for the selection conditions. The number of tuples in the result of a projection is the same as the input, assuming that duplicates are not eliminated; of course, each result tuple is smaller since it contains fewer fields.

The result size for a join can be estimated by multiplying the maximum result size, which is the product of the input table sizes, by the reduction factor of the join condition. The reduction factor for join condition *columni = column2* can be approximated by the formula $\frac{1}{MAX\ (NKeys(I1), NKeys(I2))}$ if there are indexes 11 and 12 on *columni* and *colwnn2,* respectively. This formula assumes that each key value in the smaller index, say 11, has a matching value in the other index. Given a value for *columni,* we assume that each of the *NKeys(I2)* values for *column2* is equally likely. Thus, the number of tuples that have the same value in *column2* as a given value in *columni* is $\frac{1}{NKeys(I2)}$.

12.7 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What is *metadata?* What metadata is stored in the *system catalog?* Describe the information stored per relation, and per index. (Section 12.1)
- The catalog is itself stored as a collection of relations. Explain why. (Section 12.1)
- What three techniques are commonly used in algorithms to evaluate relational operators? (Section 12.2)
- What is an access path? When does an index *match* a search condition? (Section 12.2.2)
- What are the main approaches to evaluating selections? Discuss the use of indexes, in particular. (Section 12.3.1)
- What are the main approaches to evaluating projections? What makes projections potentially expensive? (Section 12.3.2)
- What are the main approaches to evaluating joins? Why are joins expensive? (Section 12.3.3)
- What is the goal of query optimization? Is it to find the best plan? (Section 12.4)
- How does a DBMS represent a relational query evaluation plan? (Section 12.4.1)
- What is *pipelined evaluation?* What is its benefit? (Section 12.4.2)
- Describe the iterator interface for operators and access methods. 'What is its purpose? (Section 12.4.3)

- Discuss why the difference in cost between alternative plans for a query can be very large. Give specific examples to illustrate the impact of pushing selections, the choice of join methods, and the availability of appropriate indexes. (Section 12.5)
- What is the role of relational algebra equivalences in query optimization? (Section 12.6)
- What is the space of plans considered by a typical relational query optimizer? Justify the choice of this space of plans. (Section 12.6.1)
- How is the cost of a plan estimated? What is the role of the system catalog? What is the *selectivity* of an access path, and how does it influence the cost of a plan? Why is it important to be able to estimate the size of the result of a plan? (Section 12.6.2)

EXERCISES

Exercise 12.1 Briefly answer the following questions:

1. Describe three techniques commonly used when developing algorithms for relational operators. Explain how these techniques can be used to design algorithms for the selection, projection, and join operators.
2. What is an access path? When does an index *match* an access path? What is a *primary conjunct*, and why is it important?
3. What information is stored in the system catalogs?
4. What are the benefits of making the system catalogs be relations?
5. What is the goal of query optimization? Why is optimization important?
6. Describe *pipelining* and its advantages.
7. Give an example query and plan in which pipelining *cannot* be used.
8. Describe the *iterator* interface and explain its advantages.
9. What role do statistics gathered from the database play in query optimization?
10. What were the important design decisions made in the System R optimizer?
11. Why do query optimizers consider only left-deep join trees? Give an example of a query and a plan that would not be considered because of this restriction.

Exercise 12.2 Consider a relation R(a,b,c,d,e) containing 5,000,000 records, where each data page of the relation holds 10 records. R is organized as a sorted file with secondary indexes. Assume that *R.a* is a candidate key for R, with values lying in the range 0 to 4,999,999, and that R is stored in *R.o*, order. For each of the following relational algebra queries, state which of the following three approaches is most likely to be the cheapest:

- Access the sorted file for R directly.
- Use a (clustered) B+ tree index on attribute $R.a$.
- Usc a linear hashed index on attribute $R.a$.

1. $\sigma_{a<50,000}(R)$
2. $\sigma_{a=50,000}(R)$
3. $\sigma_{a>50,000 \wedge a<50,010}(R)$
4. $\sigma_{a\neq 50,000}(R)$

Exercise **12.3** For each of the following SQL queries, for each relation involved, list the attributes that must be examined to compute the answer. All queries refer to the following relations:

Emp(eid: integer, *did:* integer, *sal:* integer, *hobby:* char(20))
Dept(did: integer, *dname:* char(20), *floor:* integer, *budget:* real)

1. SELECT * FROM Emp
2. SELECT * FROM Emp, Dept
3. SELECT * FROM Emp E, Dept D WHERE E.did = D.did
4. SELECT E.eid, D,dname FROM Emp E, Dept D WHERE E.did = D.did

Exercise **12.4** Consider the following schema with the Sailors relation:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)

For each of the following indexes, list whether the index matches the given selection conditions. If there is a match, list the primary conjuncts.

1. A B+-tree index on the search key ⟨ Sailors.sid ⟩.
   (a) $\sigma_{Sailors.sid<50,000}(Sailors)$
   (b) $\sigma_{Sailors.sid=50,000}(Sailors)$
2. A hash index on the search key ⟨ Sailors.sid ⟩.
   (a) $\sigma_{Sailors.sid<50,000}(Sailors)$
   (b) $\sigma_{Sailors.sid=50,000}(Sailors)$
3. A B+-tree index on the search key ⟨ Sailors.sid, Sailors.age ⟩.
   (a) $\sigma_{Sailors.sid<50,000 \wedge Sailors.age=21}(Sailors)$
   (b) $\sigma_{Sailors.sid=50,000 \wedge Sailors.age>21}(Sailors)$
   (c) $\sigma_{Sailors.sid=50,000}(Sailors)$
   (d) $\sigma_{Sailors.age=21}(Sailors)$
4. A hash-tree index on the search key ⟨ Sailors.sid, Sailors.age ⟩.
   (a) $\sigma_{Sailors.sid=50,000 \wedge Sailors.age=21}(Sailors)$
   (b) $\sigma_{Sailors.sid=50,000 \wedge Sailors.age>21}(Sailors)$
   (c) $\sigma_{Sailors.sid=50,000}(Sailors)$
   (d) $\sigma_{Sailors.age=21}(Sailors)$

Exercise **12.5** Consider again the schema with the Sailors relation:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)

Assume that each tuple of Sailors is 50 bytes long, that a page can hold 80 Sailors tuples, and that we have 500 pages of such tuples. For each of the following selection conditions, estimate the number of pages retrieved, given the catalog information in the question.

1. Assume that we have a B+-tree index T on the search key (Sailors.sid), and assume that *IHeight(T)* = 4, *INPages(T)* = 50, *Low(T)* = 1, and *High(T)* = 100,000.
   (a) *aSailors.'id<5o,ooo(S'a'ilors)*
   (b) $\sigma_{Sailors.sid=50,000}(Sailors)$
2. Assume that we have a hash index T on the search key (Sailors.sid), and assume that *IHeight(7')* = 2, *INPages(7')* = 50, *Low(7')* = 1, and *High(T)* = 100,000.
   (a) *aSa'lor's.sid<50,OOo(Sailors)*
   (b) *aSailor·s.sid=5o,ooo(5ailors)*

Exercise **12.6** Consider the two join methods described in Section 12.3.3. Assume that we join two relations *Rand* 5, and that the systems catalog contains appropriate statistics about Rand *S.* Write formulas for the cost estimates of the index nested loops join and sort-merge join using the appropriate variables from the systems catalog in Section 12.1. For index nested loops join, consider both a B+ tree index and a hash index. (For the hash index, you can assume that you can retrieve the index page containing the rid of the matching tuple with 1.2 l/Os on average.)

Note.' Additional exercises on the material covered in this chapter can be found in the exercises for Chapters 14 and 15.

BIBLIOGRAPHIC NOTES

See the bibliograpic notes for Chapters 14 and 15.

13

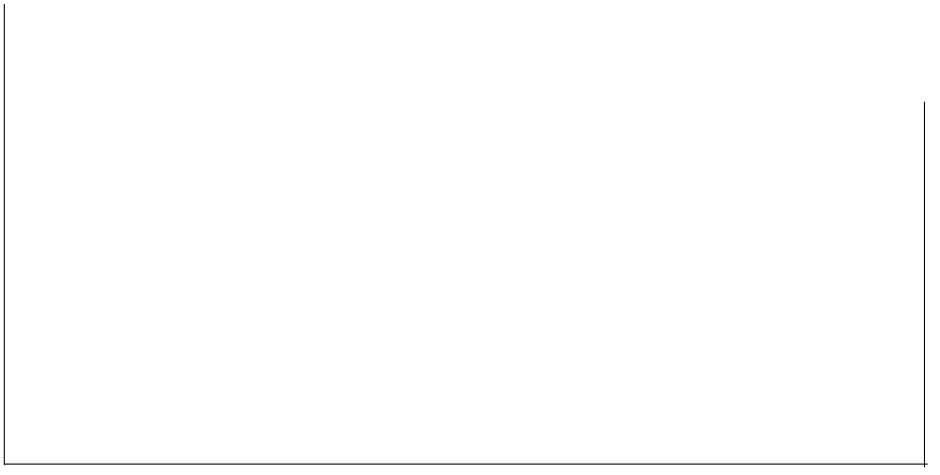
EXTERNAL SORTING

- Why is sorting important in a DBMS?
- Why is sorting data on disk different from sorting in-memory data?
- How does external merge-sort work?
- How do techniques like blockecl I/O and overlapped I/O affect the design of external sorting algorithms?
- When can we use a B+ tree to retrieve records in sorted order?
- Key concepts: motivation, bulk-loading, duplicate elimination, sort-merge joins; external merge sort, sorted runs, merging runs; replacement sorting, increasing run length; I/O cost versus number of I/Os, blocked I/Os, double buffering; B+ trees for sorting, impact of clustering.

Good order is the foundation of all things.

Edmund Burke

In this chapter, we consider a widely used and relatively expensive operation, sorting records according to a search key. We begin by considering the lnany uses of sorting in a database system in Section 13.1. We introduce the idea of external sorting by considering a very simple algorithm in Section 13.2; using repeated passes over the data, even very large datasets can be sorted with a small amount of rnemory. This algol'ithrn is generalized to develop a realistic external sorting algorithrn in Section 13.3. Three important refinements are



discussed. The first, discussed in Section 13.3.1, enables us to reduce the number of passes. The next two refinements, covered in Section 13.4, require us to consider a more detailed model of I/O costs than the number of page I/Os. Section 13.4.1 discusses the effect of *blocked* I/O, that is, reading and writing several pages at a time; and Section 13.4.2 considers how to use a technique called *double buffering* to minimize the time spent waiting for an I/O operation to complete. Section 13.5 discusses the use of B+ trees for sorting.

With the exception of Section 13.4, we consider only I/O costs, which we approximate by counting the number of pages read or written, as per the cost model discussed in Chapter 8. Our goal is to use a simple cost model to convey the main ideas, rather than to provide a detailed analysis.

13.1 WHEN DOES A DBMS SORT DATA?

Sorting a collection of records on some (search) key is a very useful operation. The key can be a single attribute or an ordered list of attributes, of course. Sorting is required in a variety of situations, including the following important ones:

- Users may want answers in some order; for example, by increasing age (Section 5.2).
- Sorting records is the first step in *bulk loading* a tree index (Section 10.8.2).
- Sorting is useful for eliminating *duplicate* copies in a collection of records (Section 14.3).

- A widely used algorithm for performing a very important relational algebra operation, called *jo'in,* requires a sorting step (Section 14.4.2).

Although main memory sizes are growing rapidly the ubiquity of database systems has lead to increasingly larger datasets as well. When the data to be sorted is too large to fit into available main memory, we need an *external sorting* algorithm. Such algorithms seek to minimize the cost of disk accesses.

13.2 A SIMPLE TWO-WAY MERGE SORT

We begin by presenting a simple algorithm to illustrate the idea behind external sorting. This algorithm utilizes only three pages of main memory, and it is presented only for pedagogical purposes. In practice, many more pages of memory are available, and we want our sorting algorithm to use the additional memory effectively; such an algorithm is presented in Section 13.3. When sorting a file, several sorted subfiles are typically generated in intermediate steps. In this chapter, we refer to each sorted subfile as a **run.**

Even if the entire file does not fit into the available main memory, we can sort it by breaking it into smaller subfiles, sorting these subfiles, and then merging them using a minimal amount of main memory at any given time. In the first pass, the pages in the file are read in one at a time. After a page is read in, the records on it are sorted and the sorted page (a sorted run one page long) is written out. Quicksort or any other in-memory sorting technique can be used to sort the records on a page. In subsequent passes, pairs of runs from the output of the previous pass are read in and *merged* to produce runs that are twice as long. This algorithm is shown in Figure 13.1.

If the number of pages in the input file is 2^k, for some k, then:

Pass 0 produces 2^k sorted runs of one page each,
Pass 1 produces 2^{k-1} sortecl runs of two pages each,
Pass 2 produces 2^{k-2} sortecl runs of four pages each,
and so on, until
Pass k produces one sorted run of 2^k pages.

In each pass, we read every page in the file, process it, and write it out. Therefore we have two disk I/Os per page, per pass. The number of passes is $\lceil log_2 N \rceil + 1$, where N is the number of pages in the file. The overall cost is $2N(\lceil log_2 N \rceil + 1)$ I/Os.

The algorithm is illustrated on all example input file containing seven pages in Figure 13.2. The sort takes four passes, and in each pass, we read and

```
proc 2-way_extsort (file)
// Given a file on disk, sorts it using three buffer pages
// Produce runs that are one page long: Pass 0
Read each page into memory, sort it, write it out.
// Merge pairs of runs to produce longer runs until only
// one run (containing all records of input file) is left
While the number of runs at end of previous pass is > 1:
    // Pass i = 1, 2, ...
    While there are runs to be merged from previous pass:
        Choose next two runs (from previous pass).
        Read each run into an input buffer; page at a time.
        Merge the runs and write to the output buffer;
        force output buffer to disk one page at a time.

endproc
```

Figure 13.1 Two-Way Merge Sort

write seven pages, for a total of 56 I/Os. This result agrees with the preceding analysis because $2 \cdot 7(\lceil \log_2 7 \rceil + 1) = 56$. The dark pages in the figure illustrate what would happen on a file of eight pages; the number of passes remains at four ($\lceil \log_2 8 \rceil + 1 = 4$), but we read and write an additional page in each pass for a total of 64 I/Os. (Try to work out what would happen on a file with, say, five pages.)

This algorithm requires just three buffer pages in main memory, as Figure 13.3 illustrates. This observation raises an important point: Even if we have more buffer space available, this simple algorithm does not utilize it effectively. The external merge sort algorithm that we discuss next addresses this problem.

13.3 EXTERNAL MERGE SORT

Suppose that B buffer pages are available in memory and that we need to sort a large file with N pages. How can we improve on the two-way merge sort presented in the previous section? The intuition behind the generalized algorithm that we now present is to retain the basic structure of making multiple passes while trying to minimize the number of passes. There are two important modifications to the two-way merge sort algorithm:

1. In Pass 0, read in B pages at a time and sort internally to produce $\lceil N/B \rceil$ runs of B pages each (except for the last run, which may contain fewer

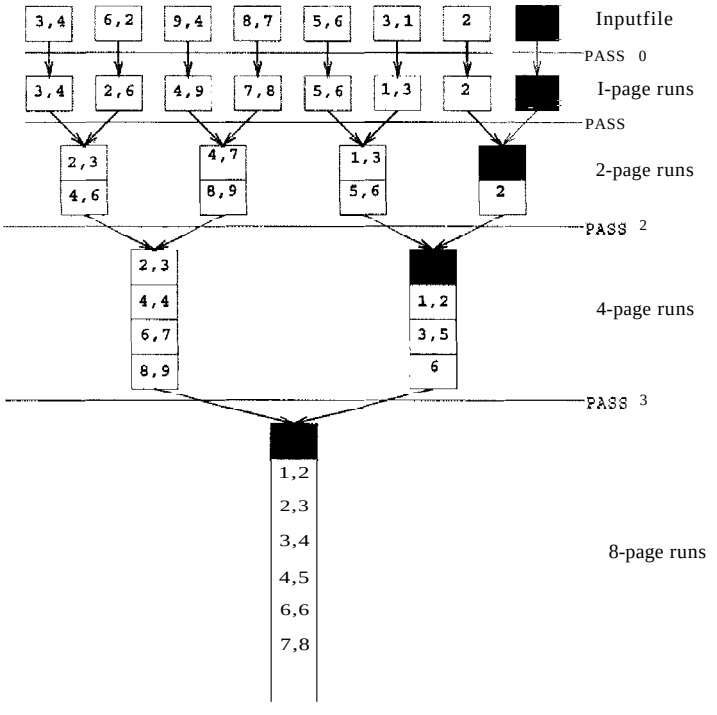


Figure 13.2 Two-Way Merge Sort of a Seven-Page File

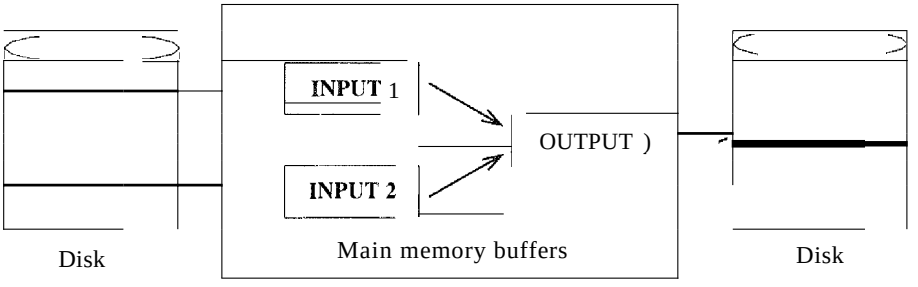


Figure 13.3 Two-'Way Merge Sort with Three Buffer Pages

pages). This modification is illustrated in Figure 13.4, using the input file from Figure 13.2 and a buffer pool with four pages.

2. In passes i = 1,2, ... use $B-1$ buffer pages for input and use the remaining page for output; hence, you do a $(B$ - I)-way merge in each pass. The utilization of buffer pages in the merging passes is illustrated in Figure 13.5.

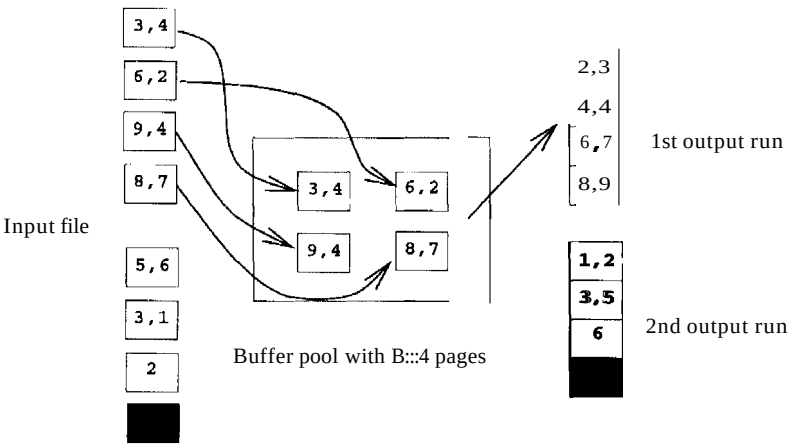


Figure 13.4 External Merge Sort with B Buffer Pages: Pass 0

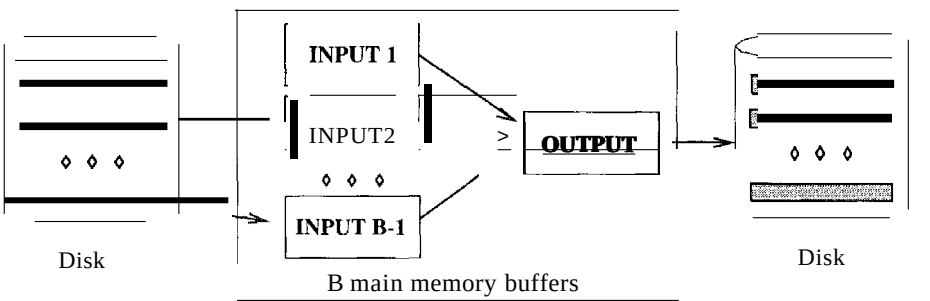


Figure 13.5 External IVlerge Sort with B Buffer Pages: Pass $i > 0$

The first refinement reduces the number of runs produced by Pass 0 to NI $\lceil N/B \rceil$, versus N for the two-way merge.[1] The second refinement is even more important. By doing a $(B-$ I)-way merge, the number of passes is reduced dramatically including the initial pass, it becomes rZ0.9B-$_1$NIl + 1 versus $\lceil log_2 N \rceil + 1$ for the two-way merge algorithm presented earlier. Because B is

[1] Note that the technique used for sorting data in buffer pages is orthogonal to external sorting. You could use, say, Quicksort for sorting data in buffer pages.

typically quite large, the savings can be substantial. The external merge sort algorithm is shown is Figure 13.6.

```
proc extsort (file)
// Given a file on disk, sorts it using three buffer pages
// Produce runs that are B pages long: Pass 0
Read B pages into memory, sort them, write out a run.
// Merge B-1 runs at a time to produce longer runs until only
// one run (containing all records of input file) is left
While the number of runs at end of previous pass is > 1:
    // Pass i = 1,2, ...
    While there are runs to be merged from previous pass:
        Choose next B -- 1 runs (from previous pass).
        Read each rUll into an input buffer; page at a time.
        Merge the rUllS and write to the output buffer;
        force output buffer to disk one page at a time.

endproc
```

Figure 13.6 External Merge Sort

As an example, suppose that we have five buffer pages available and want to sort a file with lOS pages.

Pass 0 produces $\lceil 108/5 \rceil$ = 22 sorted runs of five pages each, except for the last run, which is only three pages long.
Pass 1 does a four-way merge to produce 122/41 = six sorted runs of 20 pages each, except for the iast run, which is only eight pages long.
Pass 2 produces $\lceil 6/4 \rceil$ = two sorted runs; one with SO pages and one with 28 pages.
Pass 3 merges the two runs produced in Pass 2 to produce the sorted file.

In each pass we read and write 108 pages; thus the total cost is 2* 108*4 = 864 l/Os. Applying our formula, we have *NI* 1108/51 22 and cost $2 * N * (\lceil log_{B-1} N1 \rceil + 1)$ = $2 * 108 * (\lceil log_4 22 \rceil + 1)$ = 864, as expected.

To emphasize the potential gains in using all available buffers, in Figure 13.7, we show the number of passes, computed using our formula., for several values of Nand *B.* To obtain the cost, the number of passes should be multiplied by *2N*. In practice, one would expect to have more than 257 buffers, but this table illustrates the importance of a high fan-in during merging.

N	B = 3	B = 5	B = 9	B = 17	B = 129	B = 257
100	7	4	3	2	1	1
1000	10	5	4	3	2	2
10,000	13	7	5	4	2	2
100,000	17	9	6	5	3	3
1,000,000	20	10	7	5	3	3
10,000,000	23	12	8	6	4	3
100,000,000	26	14	9	7	4	4
1,000,000,000	30	15	10	8	5	4

Figure 13.7 Number of Passes of External Merge Sort

Of course, the CPU cost of a multiway merge can be greater than that for a two-way merge, but in general the I/O costs tend to dominate. In doing a (B - I)-way merge, we have to repeatedly pick the 'lowest' record in the $B-1$ runs being merged and write it to the output buffer. This operation can be implemented simply by examining the first (remaining) element in each of the $B-1$ input buffers. In practice, for large values of B, more sophisticated techniques can be used, although we do not discuss them here. Further, as we will see shortly, there are other ways to utilize buffer pages to reduce I/O costs; these techniques involve allocating additional pages to each input (and output) run, thereby making the number of runs merged in each pass considerably smaller than the number of buffer pages B.

13.3.1 Minimizing the Number of Runs

In Pass 0 we read in B pages at a time and sort them internally to produce $\lceil N/B \rceil$ runs of B pages each (except for the last run, which may contain fewer pages). With a more aggressive implementation, called **replacement sort,** we can write out runs of approximately $2 \cdot B$ internally sorted pages on average. This improvement is achieved as follows. We begin by reading in pages of the file of tuples to be sorted, say R, until the buffer is full, reserving (say) one page for use as an input buffer and one page for use as an output buffer. We refer to the $B - 2$ pages of R tuples that are not in the input or output buffer as the *current set.* Suppose that the file is to be sorted in ascending order on some search key k. Tuples are appended to the output in ascending order by k value.

The idea is to repeatedly pick the tuple in the current set with the smallest k value that is still greater than the largest k value in the output buffer and append it to the output buffer. For the output buffer to remain sorted, the chosen tuple must satisfy the condition that its k value be greater than or

equal to the largest k value currently in the output buffer; of all tuples in the current set that satisfy this condition, we pick the one with the smallest k value and append it to the output buffer. Moving this tuple to the output buffer creates some space in the current set, which 've use to add the next input tuple to the current set. (\Ve assume for simplicity that all tuples are the same size.) This process is illustrated in Figure 13.8. The tuple in the current set that is going to be appended to the output next is highlighted, as is the most recently appended output tuple.

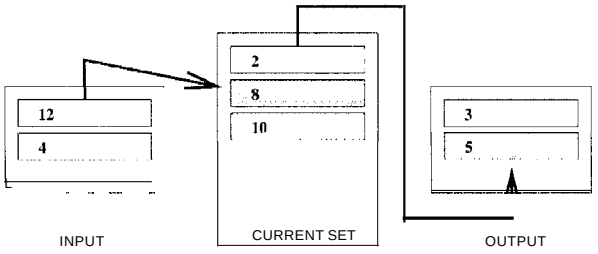


Figure 13.8 Generating Longer Runs

When all tuples in the input buffer have been consumed in this manner, the next page of the file is read in. Of course, the output buffer is written out when it is full, thereby extending the current run (which is gradually built up on disk).

The important question is this: When do we have to terminate the current run and start a new run? As long as some tuple t in the current set has a bigger k value than the most recently appended output tuple, we can append t to the output buffer and the current run can be extended.[2] In Figure 13.8, although a tuple (k = 2) in the current set has a smaller k value than the largest output tuple (k = 5), the current run can be extended because the current set also has a tuple (k = 8) that is larger than the largest output tuple.

When every tuple in the current set is smaller than the largest tuple in the output buffer, the output buffer is written out and becomes the last page in the current run. \Ve then start a new l'lm and continue the cycle of writing tuples from the input buffer to the current set to the output buffer. It is known that this algorithm produces runs that are about $2 \cdot B$ pages long, on average.

This refinement has not been implemented in commercial database systenls because managing the main memory available for sorting becOlnes difficult with

2If B is large, the CPU cost of finding such a tuple t can be significant unless appropriate in-memory data structures are used to organize the tuples in the buffer pool. \Ve will not discuss this issue further.

replacement sort, especially in the presence of variable length records. Recent work on this issue, however, shows promise and it could lead to the use of replacement sort in commercial systems.

13.4 MINIMIZING I/O COST VERSUS NUMBER OF I/OS

We have thus far used the number of page 1/Os as a cost metric. This metric is only an approximation of true I/O costs because it ignores the effect of *blocked* I/O--issuing a single request to read (or write) several consecutive pages can be much cheaper than reading (or writing) the same number of pages through independent I/O requests, as discussed in Chapter 8. This difference turns out to have some very important consequences for our external sorting algorithm.

Further, the time taken to perform I/O is only part of the time taken by the algorithm; we must consider CPU costs as well. Even if the time taken to do I/O accounts for most of the total time, the time taken for processing records is nontrivial and definitely worth reducing. In particular, we can use a technique called *double buffeTing* to keep the CPU busy while an I/O operation is in progress.

In this section, we consider how the external sorting algorithm can be refined using blocked I/O and double buffering. The motivation for these optimizations requires us to look beyond the number of I/Os as a cost metric. These optimizations can also be applied to other I/O intensive operations such as joins, which we study in Chapter 14.

13.4.1 Blocked I/O

If the number of page I/Os is taken to be the cost metric, the goal is clearly to minimize the number of passes in the sorting algorithm because each page in the file is read and written in each pass. It therefore makes sense to maximize the fan-in during merging by allocating just one buffer pool page per run (which is to be merged) and one buffer page for the output of the merge. Thus, we can merge $B-1$ runs, where B is the number of pages in the buffer pool. If we take into account the effect of blocked access, which reduces the average cost to read or write *a single page*, we are led to consider whether it might be better to read and write in units of more than one page.

Suppose we decide to read and write in units, which we call **buffer** blocks, of b pages. We must now set aside one buffer block per input run and one buffer block for the output of the merge, which means that we can merge at most $\lfloor \frac{B-b}{b} \rfloor$ runs in each pass. For example, if we have 10 buffer pages, we can either merge nine runs at a time with one-page input and output buffer

blocks, or we can merge four runs at a time with two-page input and output buffer blocks. If we choose larger buffer blocks, however, the number of passes increases, while we continue to read and write every page in the file in each pass! In the example, each merging pass reduces the number of runs by a factor of 4, rather than a factor of 9. Therefore, the number of page I/Os increases. This is the price we pay for decreasing the per-page I/O cost and is a trade-off we must take into account when designing an external sorting algorithm.

In practice, however, current main memory sizes are large enough that all but the largest files can be sorted in just two passes, even using blocked I/O. Suppose we have B buffer pages and choose to use a blocking factor of b pages. That is, we read and write b pages at a time, and all our input and output buffer blocks are b pages long. The first pass produces about $N2 = \lceil N/2B \rceil$ sorted runs, each of length $2B$ pages, if we use the optimization described in Section 13.3.1, and about $N1 = \lceil N/B \rceil$ sorted runs, each of length B pages, otherwise. For the purposes of this section, we assume that the optimization is used.

In subsequent passes we can merge $F = \lfloor B/b \rfloor - 1$ runs at a time. The number of passes is therefore $1 + \lceil \log_F N2 \rceil$, and in each pass we read and write all pages in the file. Figure 13.9 shows the number of passes needed to sort files of various sizes N, given B buffer pages, using a blocking factor b of 32 pages. It is quite reasonable to expect 5000 pages to be available for sorting purposes; with 4KB pages, 5000 pages is only 20MB. (With 50,000 buffer pages, we can do 1561-way merges; with 10,000 buffer pages, we can do 311-way merges; with 5000 buffer pages, we can do 155-way merges; and with 1000 buffer pages, we can do 30-way merges.)

N	B = 1000	B = 5000	B = 10,000	B = 50,000
100	1	1	1	1
1000	1	1	1	1
10,000	2	2	1	1
100,000	3	2	2	2
1,000,000	3	2	2	2
10,000,000	4	3	3	2
100,000,000	5	3	3	2
1,000,000,000	5	4	3	3

Figure 13.9 Number of Passes of External Merge Sort with Block Size $b = 32$

To compute the I/O cost, we need to calculate the number of 32-page blocks read or written and multiply this number by the cost of doing a 32-page block I/O. To find the number of block I/Os, we can find the total number of page

l/Os (number of passes rnultiplied by the number of pages in the file) and divide by the block size, 32. The cost of a 32-page block I/O is the seek time and rotational delay for the first page, plus transfer time for all 32 pages, as discussed in Chapter 8. The reader is invited to calculate the total I/O cost of sorting files of the sizes mentioned in Figure 13.9 with 5000 buffer pages for different block sizes (say, b = 1, 32, and 64) to get a feel for the benefits of using blocked I/O.

13.4.2 Double Buffering

Consider what happens in the external sorting algorithm when all the tuples in an input block have been consumed: An I/O request is issued for the next block of tuples in the corresponding input run, and the execution is forced to suspend until the I/O is complete. That is, for the duration of the time taken for reading in one block, the CPU remains idle (assuming that no other jobs are running). The overall time taken by an algorithm can be increased considerably because the CPU is repeatedly forced to wait for an I/O operation to complete. This effect becomes more and more important as CPU speeds increase relative to I/O speeds, which is a long-standing trend in relative speeds. It is therefore desirable to keep the CPU busy while an I/O request is being carried out; that is, to overlap CPU and I/O processing. Current hardware supports such overlapped computation, and it is therefore desirable to design algorithms to take advantage of this capability.

In the context of external sorting, we can achieve this overlap by allocating extra pages to each input buffer. Suppose a block size of b = 32 is chosen. The idea is to allocate an additional 32-page block to every input (and the output) buffer. Now, when all the tuples in a 32-page block have been consumed, the CPU can process the next 32 pages of the run by switching to the second, 'double,' block for this run. Meanwhile, an I/O request is issued to fill the empty block. Thus, assmning that the time to consume a block is greater than the time to read in a block, the CPU is never idle! On the other hand, the number of pages allocated to a buffer is doubled (for a given block size, which means the total I/O cost stays the same). This technique, called double buffering, can considerably reduce the total time taken to sort a file. The use of buffer pages is illustrated in Figure 13.10.

Note that although double buffering can considerably reduce the response tiule for a given query, it may not have a significant impact on throughput, because the CPU can be kept busy by working on other queries while waiting for one query's I/O operation to complete.

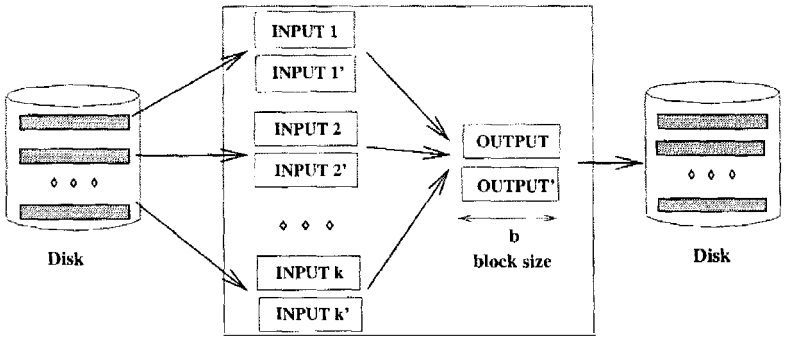


Figure 13.10 Double Buffering

13.5 USING B+ TREES FOR SORTING

Suppose that we have a B+ tree index on the (search) key to be used for sorting a file of records. Instead of using an external sorting algorithm, we could use the B+ tree index to retrieve the records in search key order by traversing the sequence set (i.e., the sequence of leaf pages). Whether this is a good strategy depends on the nature of the index.

13.5.1 Clustered Index

If the B+ tree index is clustered, then the traversal of the sequence set is very efficient. The search key order corresponds to the order in which the data records are stored, and for each page of data records we retrieve, we can read all the records on it in sequence. This correspondence between search key ordering and data record ordering is illustrated in Figure 13.11, with the assumption that data entries are *(key, rid)* pairs (i.e., Alternative (2) is used for data entries).

The cost of using the clustered B+ tree index to retrieve the data records in search key order is the cost to traverse the tree from root to the left-most leaf (which is usually less than four I/Os) plus the cost of retrieving the pages in the sequence set, plus the cost of retrieving the (say, *N)* pages containing the data records. Note that no data page is retrieved twice, thanks to the ordering of data entries being the same as the ordering of data records. The number of pages in the sequence set is likely to be much smaller than the number of data pages because data entries are likely to be smaller than typical data records. Thus, the strategy of using a dusterecl B+ tree inclex to retrieve the records in sorted order is a good one and should be used whenever such an index is ailable.

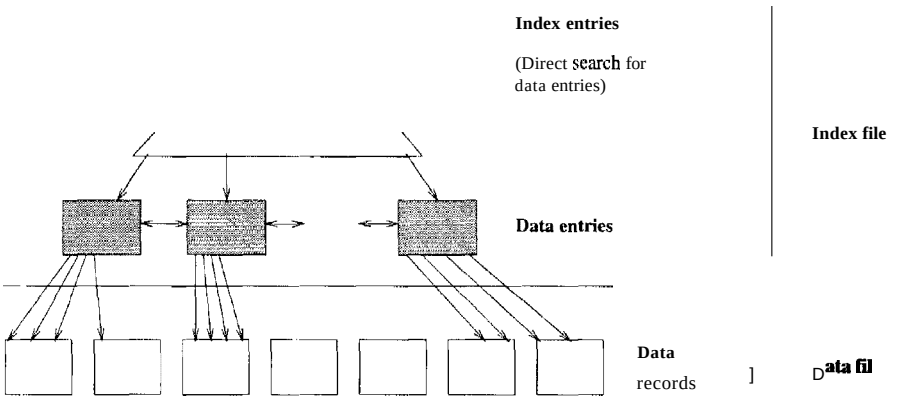


Figure 13.11 Clustered B+ Tree for Sorting

What if Alternative (1) is used for data entries? Then, the leaf pages would contain the actual data records, and retrieving the pages in the sequence set (a total of N pages) would be the only cost. (Note that the space utilization is about 67% in a B+ tree; the number of leaf pages is greater than the number of pages needed to hold the data records in a sorted file, where, in principle, 100% space utilization can be achieved.) In this case, the choice of the B+ tree for sorting is excellent!

13.5.2 Unclustered Index

What if the B+ tree index on the key to be used for sorting is unclustered? This is illustrated in Figure 13.12, with the assumption that data entries are $\langle key, rid \rangle$.

In this case each rid in a leaf page could point to a different data page. Should this happen, the cost (in disk l/Os) of retrieving all data records could equal the number of data records. That is, the worst-case cost is equal to the number of data records, because fetching each record could require a disk I/O. This cost is in addition to the cost of retrieving leaf pages of the B+ tree to get the data entries (which point to the data records).

If p is the average number of records per data page and there are N data pages, the number of data records is $p \cdot N$. If we take f to be the ratio of the size of a data entry to the size of a data record, we can approximate the number of leaf pages in the tree by $f \cdot N$. The total cost of retrieving records in sorted order

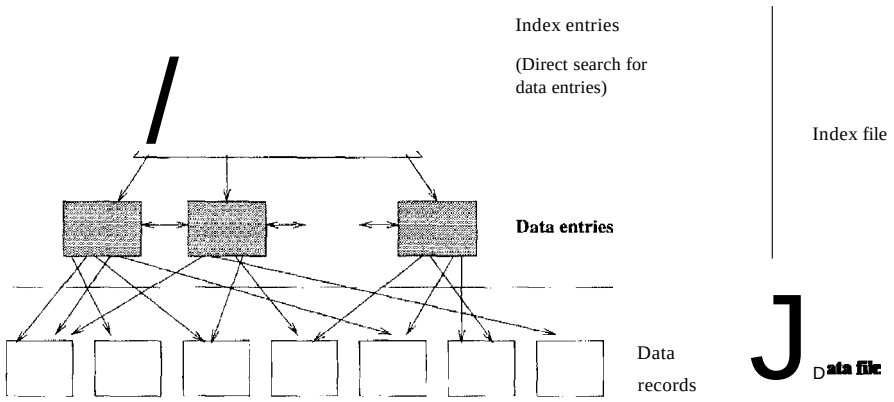


Figure 13.12 Unclustered B+ Tree for Sorting

using an unclustered B+ tree is therefore $(J + p) . N$. Since f is usually 0.1 or smaller and p is typically much larger than 10, $p . N$ is a good approximation.

In practice, the cost may be somewhat less because some rids in a leaf page lead to the same data page, and further, some pages are found in the buffer pool, thereby avoiding an I/O. Nonetheless, the usefulness of an unclustered B+ tree index for sorted retrieval highly depends on the extent to which the order of data entries corresponds and--this is just a matter of chance-to the physical ordering of data records.

We illustrate the cost of sorting a file of records using external sorting and unclustered B+ tree indexes in Figure 13.13. The costs shown for the unclustered index are worst-case numbers, based on the approximate formula $p . N$. For comparison, note that the cost for a clustered index is approximately equal to N, the number of pages of data records.

	Sorting	$p = 1$	$p = 10$	$p = 100$
100	200	100	1000	10,000
1000	2000	1000	10,000	100,000
10,000	40,000	10,000	100,000	1,000,000
100,000	600,000	100,000	1,000,000	10,000,000
1,000,000	8,000,000	1,000,000	10,000,000	100,000,000
10,000,000	80,000,000	10,000,000	100,000,000	1,000,000,000

Figure 13.13 Cost of External Sorting ($13 = 1000$, $b = 32$) versus Unclustered Index

Keep in mind that p is likely to be doser to 100 and B is likely to be higher than 1,000 in practice. The ratio of the cost of sorting versus the cost of using an unclustered index is likely to be even lower than indicated by Figure 13.13 because the I/O for sorting is in 32-page buffer blocks, whereas the I/O for the undustered indexes is one page at a time. The value of p is determined by the page size and the size of a data record; for p to be 10, with 4KB pages, the average data record size must be about 400 bytes. In practice, p is likely to be greater than 10.

For even modest file sizes, therefore, sorting by using an unclustered index is clearly inferior to external sorting. Indeed, even if we want to retrieve only about 10--20% of the data records, for example, in response to a range query such as "Find all sailors whose rating is greater than 7," sorting the file may prove to be more efficient than using an unclustered index!

13.6 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What database operations utilize sorting? (Section 13.1)
- Describe how the *two-way merge sort* algorithm can sort a file of arbitrary length using only three main-memory pages at any time. Explain what a *run* is and how runs are created and merged. Discuss the cost of the algorithm in terms of the number of *passes* and the I/O cost per pass. (Section 13.2)
- How does the general *external merge sort* algorithm improve upon the two-way merge sort? Discuss the length of initial runs, and how memory is utilized in subsequent merging passes. Discuss the cost of the algorithm in terms of the number of passes and the I/O cost per pass. (Section 13.3)
- Discuss the use of *replacement sort* to increase the average length of initial runs and thereby reduce the number of runs to be merged. How does this affect the cost of external sorting? (Section 13.3.1)
- What is *blocked I/O?* Why is it cheaper to read a sequence of pages using blocked I/O than to read them through several independent requests? How does the use of blocking affect the external sorting algorithm, and how does it change the cost formula? (Section 13.4.1)
- What is *double buffering?* What is the motivation for using it? (Section 13.4.2)
- If we want to sort a file and there is a B+ tree with the same search key, we have the option of retrieving records in order through the index. Compare

the cost of this approach to retrieving the records in random order and then sorting them. Consider both clustered and unclustered B+ trees. What conclusions can you draw from your comparison? **(Section 13.5)**

EXERCISES

Exercise 13.1 Suppose you have a file with 10,000 pages and you have three buffer pages. Answer the following questions for each of these scenarios, assuming that our most general external sorting algorithm is used:

(a) A file with 10,000 pages and three available buffer pages.

(b) A file with 20,000 pages and five available buffer pages.

(c) A file with 2,000,000 pages and 17 available buffer pages.

1. How many runs will you produce in the first pass?
2. How many passes will it take to sort the file completely?
3. What is the total I/O cost of sorting the file?
4. How many buffer pages do you need to sort the file completely in just two passes?

Exercise 13.2 Answer Exercise 13.1 assuming that a two-way external sort is used.

Exercise 13.3 Suppose that you just finished inserting several records into a heap file and now want to sort those records. Assume that the DBMS uses external sort and makes efficient use of the available buffer space when it sorts a file. Here is some potentially useful information about the newly loaded file and the DBMS software available to operate on it:

> The number of records in the file is 4500. The sort key for the file is 4 bytes long. You can assume that rids are 8 bytes long and page ids are 4 bytes long. Each record is a total of 48 bytes long. The page size is 512 bytes. Each page has 12 bytes of control information on it. Four buffer pages are available.

1. How many sorted subfiles will there be after the initial pass of the sort, and how long will each subtile be?
2. How many passes (including the initial pass just considered) are required to sort this file?
3. What is the total I/O cost for sorting this file?
4. What is the largest file, in terms of the number of records, you can sort with just four buffer pages in two passes? How would your answer change if you had 257 buffer pages?
5. Suppose that you have a B+ tree index with the search key being the same as the desired sort key. Find the cost of using the index to retrieve the records in sorted order for each of the following cases:
   - The index uses Alternative (1) for data entries.
   - The index uses Alternative (2) and is unclustered. (You can compute the worst-case cost in this case.)

- How would the costs of using the index change if the file is the largest that you can sort in two passes of external sort with 257 buffer pages? Give your answer for both clustered and unclustered indexes.

Exercise 13.4 Consider a disk with an average seek time of lOms, average rotational delay of 5ms, and a transfer time of 1ms for a 41< page. Assume that the cost of reading/writing a page is the sum of these values (i.e., 16ms) unless a *sequence* of pages is read/written. In this case, the cost is the average seek time plus the average rotational delay (to find the first page in the sequence) plus 1ms per page (to transfer data). You are given 320 buffer pages and asked to sort a file with 10,000,000 pages.

1. Why is it a bad idea to use the 320 pages to support virtual memory, that is, to 'new' 10,000,000 4K bytes of memory, and to use an in-memory sorting algorithm such as Quicksort?
2. Assume that you begin by creating sorted runs of 320 pages each in the first pass. Evaluate the cost of the following approaches for the subsequent merging passes:
   (a) Do 31g-way merges.
   (b) Create 256 'input' buffers of 1 page each, create an 'output' buffer of 64 pages, and do 256-way merges.
   (c) Create 16 'input' buffers of 16 pages each, create an 'output' buffer of 64 pages, and do 16-way merges.
   (d) Create eight 'input' buffers of 32 pages each, create an 'output' buffer of 64 pages, and do eight-way merges.
   (e) Create four 'input' buffers of 64 pages each, create an 'output' buffer of 64 pages, and do four-way merges.

Exercise 13.5 Consider the refinement to the external sort algorithm that produces runs of length *2B* on average, where *B* is the number of buffer pages. This refinement was described in Section 11.2.1 under the assumption that all records are the same size. Explain why this assumption is required and extend the idea to cover the case of variable-length records.

PROJECT-BASED EXERCISES

Exercise 13.6 *(Note to instructors: Additional deta'ils must be provided if this exercise* is *assigned; see Appendix 30.)* Implement external sorting in Minibase.

BIBLIOGRAPHIC NOTES

Knuth's text [442] is the classic reference for sorting algorithms. Memory management for replacement sort is discussed in [471]. A number of papers discuss parallel external sorting algorithms, including [66, 71, 223, 494, 566, 647].

14

EVALUATING RELATIONAL OPERATORS

- What are the alternative algorithms for selection? Which alternatives are best under different conditions? How are complex selection conditions handled?
- How can we eliminate duplicates in projection? How do sorting and hashing approaches compare?
- What are the alternative join evaluation algorithms? Which alternatives are best under different conditions?
- How are the set operations (union, inter;section, set-difference, cross-product) implemented?
- How are aggregate operations and grouping handled?
- How does the size of the buffer pool and the buffer replacement policy affect algorithms for evaluating relational operators?
- **Key concepts:** selections, CNF; projections, sorting versus hashing; joins, block nested loops, index nested loops, sort-merge, hash; union, set-difference, duplicate elimination; aggregate operations, running information, partitioning into groups, using indexes; buffer management, concurrent execution, repeated access patterns

Now, *here*, you see, it takes all the running you can do, to keep in the same place. If you want to get somewhere else, you must run at least twice as fast as that!

—Lewis Carroll, *Through the Looking Glass*

In this chapter, we consider the implementation of individual relational operators in sufficient detail to understand how DBMSs are implemented. The discussion builds on the foundation laid in Chapter 12. We present implementation alternatives for the selection operator in Sections 14.1 and 14.2. It is instructive to see the variety of alternatives and the wide variation in per'formanee of these alternatives, for even such a simple operator. In Section 14.3, we consider the other unary operator in relational algebra, projection.

We then discuss the implementation of binary operators, beginning with joins in Section 14.4. Joins are among the most expensive operators in a relational database system, and their implementation has a big impact on performance. After discussing the join operator, we consider implementation of the binary operators cross-product, intersection, union, and set-difference in Section 14.5. We discuss the implementation of grouping and aggregate operators, which are extensions of relational algebra, in Section 14.6. We conclude with a discussion of how buffer management affects operator evaluation costs in Section 14.7.

The discussion of each operator is largely independent of the discussion of other operators. Several alternative implementation techniques are presented for each operator; the reader who wishes to cover this material ill less depth can skip some of these alternatives without loss of continuity.

Preliminaries: Examples and Cost Calculations

We present a number of example queries using the same schema as in Chapter 12:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)
Reserves(*sid:* integer, *bid:* integer, *day:* dates, *rname:* string)

This schema is a variant of the one that we used in Chapter 5; we added a string field *rname* to Reserves. Intuitively, this field is the name of the person who made the reservation (and may be different from the name of the sailor *.sid* for whom the reservation was made; a reservation may be made by a person who is not a sailor on behalf of a sailor). The addition of this field gives us more flexibility in choosing illustrative examples. We assume that each tuple of Reserves is 40 bytes lOllg, that a page can hold 100 Reserves tuples, alld that we have 1000 pages of such tuples. Similarly, we assume that each tuple of Sailors is 50 bytes long, that a page can hold 80 Sailors tuples, and that we have 500 pages of such tuples.

Two points must be kept in Inind to understancl our discussion of costs:

- As discussed in Chapter 8, we consider only I/O costs and measure I/O cost in terms of the number of page l/Os. We also use big-O notation to express the complexity of an algorithm in terms of an input parameter and assume that the reader is familiar with this notation. For example, the cost of a file scan is $O(M)$, where M is the size of the file.
- We discuss several alternate algorithms for each operation. Since each alternative incurs the same cost in writing out the result, should this be necessary, we uniformly ignore this cost in comparing alternatives.

14.1 THE SELECTION OPERATION

In this section, we describe various algorithms to evaluate the selection operator. To motivate the discussion, consider the selection query shown in Figure 14.1, which has the selection condition *rno:me='Joe'*.

```
SELECT *
FROM   Reserves R
WHERE  R.rname='Joe'
```

Figure 14.1 Simple Selection Query

We can evaluate this query by scanning the entire relation, checking the condition on each tuple, and adding the tuple to the result if the condition is satisfied. The cost of this approach is 1000 l/Os, since Reserves contains 1000 pages. If only a few tuples have *rnarne= 'Joe',* this approach is expensive because it does not utilize the selection to reduce the number of tuples retrieved in any way. How can we improve on this approach? The key is to utilize information in the selection condition and use an index if a suitable index is available. For example, a B+ tree index on *rname* could be used to answer this query considerably faster, but an index on *bid* would not be useful.

In the rest of this section. we consider various situations with respect to the file organization used for the relation and the availability of indexes and discuss appropriate algorithms for the selection operation. We discuss only simple selection operations of the form $\sigma_{R.attr}$ **op** $_{value}(R)$ until Section 14.2, where we consider general selections. In terms of the general techniques listed in Section 12.2, the algorithms for selection use either iteration or indexing.

14.1.1 No Index, Unsorted Data

Given a selection of the form $\sigma_{R.attr}$ **op** $_{value}(R)$, if there is no index on $R.attr$ and R is not sorted on $R.attr$, we have to scan the entire relation. Therefore,

the most selective access path is a file scan. For each tuple, we must test the condition *R.attr* **op** *value* and add the tuple to the result if the condition is satisfied.

14.1.2 No Index, Sorted Data

Given a selection of the form *O'R.attr* **op** *value(R),* if there is no index on *R.attr,* but *R* is physically sorted on *R.attr,* we can utilize the sort order by doing a binary search to locate the first tuple that satisfies the selection condition. Further, we can then retrieve all tuples that satisfy the selection condition by starting at this location and scanning *R* until the selection condition is no longer satisfied. The access method in this case is a sorted-file scan with selection condition $\sigma_{R.attr}$ **op** *value(R).*

For example, suppose that the selection condition is *R.aUr!* > 5, and that *R* is sorted on *attr1* in ascending order. After a binary search to locate the position in *R* corresponding to 5, we simply scan all remaining records.

The cost of the binary search is *O(l092M).* In addition, we have the cost of the scan to retrieve qualifying tuples. The cost of the scan depends on the number of such tuples and can vary from zero to M. In our selection from Reserves (Figure 14.1), the cost of the binary search is *[0921000* $\approx$ 10 I/Os.

In practice, it is unlikely that a relation will be kept sorted if the DBMS supports Alternative (1) for index data entries; that is, allows data records to be stored as index data entries. If the ordering of data records is important, a better way to maintain it is through a B+ tree index that uses Alternative (1).

14.1.3 B+ Tree Index

If a clustereel B+ tree index is available on *R.attr,* the best strategy for selection conditions $\sigma_{R.attr}$ **op** *value(R)* in which **op** is not equality is to use the index. This strategy is also a good access path for equality selections, although a hash index on *R.attr* would be a little better. If the B+ tree index is not clustered, the cost of using the index depends on the number of tuples that satisfy the selection, as discussed later.

We can use the index as follows: We search the tree to find the first index entry that points to a qualifying tuple of *R.* Then we scan the leaf pages of the index to retrieve all entries in which the key value satisfies the selection condition. For each of these entries, we retrieve the corresponding tuple of *R.* (For concreteness in this discussion, we assume that data entries use Alternatives (2) or (3); if Alternative (1) is used, the data entry contains the actual tuple

and there is no additional cost—beyond the cost of retrieving data entries—for retrieving tuples.)

The cost of identifying the starting leaf page for the scan is typically two or three I/Os. The cost of scanning the leaf level page for qualifying data entries depends on the number of such entries. The cost of retrieving qualifying tuples from *R* depends on two factors:

- The number of qualifying tuples.
- Whether the index is clustered. (Clustered and unclustered B+ tree indexes are illustrated in Figures 13.11 and 13.12. The figures should give the reader a feel for the impact of clustering, regardless of the type of index involved.)

If the index is clustered, the cost of retrieving qualifying tuples is probably just one page I/O (since it is likely that all such tuples are contained in a single page). If the index is not clustered, each index entry could point to a qualifying tuple on a different page, and the cost of retrieving qualifying tuples in a straightforward way could be one page I/O per qualifying tuple (unless we get lucky with buffering). We can significantly reduce the number of I/Os to retrieve qualifying tuples from *R* by first sorting the rids (in the index's data entries) by their *page-id* component. This sort ensures that, when we bring in a page of *R,* all qualifying tuples on this page are retrieved one after the other. The cost of retrieving qualifying tuples is now the number of pages of *R* that contain qualifying tuples.

Consider a selection of the form *rnarne* < *'C%'* on the Reserves relation. Assuming that names are uniformly distributed with respect to the initial letter, for simplicity, we estimate that roughly 10% of Reserves tuples are in the result. This is a total of 10,000 tuples, or 100 pages. If we have a clustered B+ tree index on the *marne* field of Reserves, we can retrieve the qualifying tuples with 100 I/Os (plus a few I/Os to traverse from the root to the appropriate leaf page to start the scan). However, if the index is unclustered, we could have up to 10,000 I/Os in the worst case, since each tuple could cause us to read a page. If we sort the rids of Reserves tuples by the page number and then retrieve pages of Reserves, we avoid retrieving the same page multiple times; nonetheless, the tuples to be retrieved are likely to be scattered across many more than 100 pages. Therefc)re, the use of an unclusterecl index for a range selection could be expensive; it might be cheaper to simply scan the entire relation (which is lOOn pages in our example).

14.1.4 Hash Index, Equality Selection

If a hash index is available on $R.attr$ and **op** is equality, the best way to implement the selection $_{CTR.attr}\,\mathbf{op}\,_{value}(R)$ is obviously to use the index to retrieve qualifying tuples.

The cost includes a few (typically one or two) I/Os to retrieve the appropriate bucket page in the index, plus the cost of retrieving qualifying tuples from R. The cost of retrieving qualifying tuples from R depends on the number of such tuples and on whether the index is clustered. Since **op** is equality, there is exactly one qualifying tuple if $R.attr$ is a (candidate) key for the relation. Otherwise, we could have several tuples with the same value in this attribute.

Consider the selection in Figure 14.1. Suppose that there is an unclustered hash index on the *rname* attribute, that we have 10 buffer pages, and that 100 reservations were made by people named Joe. The cost of retrieving the index page containing the rids of such reservations is one or two I/Os. The cost of retrieving the 100 Reserves tuples can vary between 1 and 100, depending on how these records are distributed across pages of Reserves and the order in which we retrieve these records. If these 100 records are contained in, say, some five pages of Reserves, we have just five additional I/Os if we sort the rids by their page component. Otherwise, it is possible that we bring in one of these five pages, then look at some of the other pages, and find that the first page has been paged out when we need it again. (Remember that several users and DBMS operations share the buffer pool.) This situation could cause us to retrieve the same page several times.

14.2 GENERAL SELECTION CONDITIONS

In our discussion of the selection operation thus far, we have considered selection conditions of the form $_{CTR.attr}\,\mathbf{op}\,_{value(R)}$. In general, a selection condition is a Boolean combination (Le., an expression using the logical connectives $1\backslash$ and V) of terms that have the form *attribute* **op** *constant* or *attributel* **op** *attrilmte2*. For example, if the WHERE clause in the query shown in Figure 14.1 contained the condition *R.rnarne='Joe'* AND *R.bid=r*, the equivalent algebra expression would be $_{CTR.rname='Joe'l\backslash R.bid=r(R)}$.

In Section 14.2.1, we provide a more rigorous definition of CNF, which we introduced in Section 12.2.2. We consider algorithms for applying selection conditions without disjunction in Section 14.2.2 and then discuss conditions with disjunction in Section 14.2.3.

14.2.1 CNF and Index Matching

To process a selection operation with a general selection condition, we first express the condition in conjunctive normal form (CNF), that is, as a collection of *conjunets* that are connected through the use of the $\wedge$ operator. Each conjunct consists of one or more *terms* (of the form described previously) connected by V.[1] Conjuncts that contain V are said to be disjunctive or to contain disjunction.

As an example, suppose that we have a selection on Reserves with the condition *(day* < *8/9/02* $\wedge$ *rname* = *'Joe'*) $\vee$ *bid=5* $\vee$ *sid=3.* We can rewrite this in conjunctive normal form as *(day* < *8/9/02* $\vee$ *bid=5* $\vee$ *sid=3)* $\wedge$ *(rname* = *'Joe'V bid=5* $\vee$ *sid=3).*

We discussed when an index matches a CNF selection in Section 12.2.2 and introduced selectivity of access paths. The reader is urged to review that material now.

14.2.2 Evaluating Selections without Disjunction

When the selection does not contain disjunction, that is, it is a conjunction of terms, we have two evaluation options to consider:

- We can retrieve tuples using a file scan or a single index that matches some conjuncts (and which we estimate to be the most selective access path) and apply all nonprimary conjuncts in the selection to each retrieved tuple. This approach is very similar to how we use indexes for simple selection conditions, and we do not discuss it further. (We emphasize that the number of tuples retrieved depends on the selectivity of the primary conjuncts in the selection, and the remaining conjuncts only reduce the cardinality of the result of the selection.)
- We can try to utilize several indexes. We examine this approach in the rest of this section.

If several indexes containing data entries with rids (i.e., Alternatives (2) or (3)) match conjuncts in the selection, we can use these indexes to compute sets of rids of candidate tuples. We can then intersect these sets of rids, typically by first sorting them, then retrieving those records whose rids are in the intersection. If additional conjuncts are present in the selection, we can apply these conjuncts to discard some of the candidate tuples from the result.

[1] Every selection conditioll can be expressed in CNF. We refer the reader to any standard text on mathematical logic for the details.

Intersecting rid Sets: Oracle 8 uses several techniques to do rid set intersection for selections with AND. One is to AND bit maps. Another is to do a hash join of indexes. For example, given $sal < 5 \wedge price > 30$ and indexes on *sal* and *price*, we can join the indexes on the rid column, considering only entries that satisfy the given selection conditions. Microsoft SQL Server implements rid set intersection through index joins. IBM DB2 implements intersection of rid sets using Bloom filters (which are discussed in Section 22.10.2). Sybase ASE does not do rid set intersection for AND selections; Sybase ASIQ does it using bitmap operations. Informix also does rid set intersection.

As an example, given the condition *day* < *8/9/02* ∧ *bid=5* ∧ *sid=,'J,* we can retrieve the rids of records that meet the condition *day* < *8/9/02* by using a B+ tree index on *day,* retrieve the rids of records that meet the condition *sid=,'J* by using a hash index on *sid,* and intersect these two sets of rids. (If we sort these sets by the page id component to do the intersection, a side benefit is that the rids in the intersection are obtained in sorted order by the pages that contain the corresponding tuples, which ensures that we do not fetch the same page twice while retrieving tuples using their rids.) We can now retrieve the necessary pages of Reserves to retrieve tuples and check *bid=5* to obtain tuples that meet the condition *day* < *8/9/02* ∧ *bid=5* ∧ *sid=,'J.*

14.2.3 Selections with Disjunction

Now let us consider that one of the conjuncts in the selection condition is a *disjunction of terms.* If even one of these terms requires a file scan because suitable indexes or sort orders are unavailable, testing this conjunct by itself (I.e., without taking advantage of other conjuncts) requires a file scan. For example, suppose that the only available indexes are a hash index on *rname* and a hash index on *sid,* and that the selection condition contains just the (disjunctive) conjunct *(day* < *8/9/02* V *rnarne='Joe').* We can retrieve tuples satisfying the condition *rname='Joe'* by using the index on *rnarne.* However, *day* < *8/9/02* requires a file scan. So we might as well do a file scan and check the condition *rname='Joe'* for each retrieved tuple. Therefore, the most selective access path in this example is a file scan.

On the other hand, if the selection condition is *(day* < *8/9/02* ∨ *rname='Joe')* ∧ *sid=,'J,* the index on *sid* matches the conjunct *sid=S.* We can use this index to find qualifying tuples and apply *day* < *8/9/02* ∨ *rname='Joe'* to just these tuples. The best access path in this example is the index on *sid* with the primary conjunct *sid=3.*

Disjunctions: Microsoft SQL Server considers the use of unions and bitmaps for dealing with disjunctive conditions. Oracle 8 considers four ways to handle disjunctive conditions: (1) Convert the query into a union of queries without OR. (2) If the conditions involve the same attribute, such as $sal < 5 \vee sal > 30$, use a nested query with an IN list and an index on the attribute to retrieve tuples matching a value in the list. (3) Use bitmap operations, e.g., evaluate $sal < 5 \vee sal > 30$ by generating bitmaps for the values 5 and 30 and OR the bitmaps to find the tuples that satisfy one of the conditions. (We discuss bitmaps in Chapter 25.) (4) Simply apply the disjunctive condition as a filter on the set of retrieved tuples. Sybase ASE considers the use of unions for dealing with disjunctive queries and Sybase ASIQ uses bitmap operations.

Finally, if every term in a disjunction has a matching index, we can retrieve candidate tuples using the indexes and then take the union. For example, if the selection condition is the conjunct $(day < 8/9/02 \vee rname='Joe')$ and we have B+ tree indexes on *day* and *rname*, we can retrieve all tuples such that $day < 8/9/02$ using the index on *day*, retrieve all tuples such that $rname='Joe'$ using the index on *rname*, and then take the union of the retrieved tuples. If all the matching indexes use Alternative (2) or (3) for data entries, a better approach is to take the union of rids and sort them before retrieving the qualifying data records. Thus, in the example, we can find rids of tuples such that $day < 8/9/02$ using the index on *day*, find rids of tuples such that $rname='Joe'$ using the index on *rname*, take the union of these sets of rids and sort them by page number, and then retrieve the actual tuples from Reserves. This strategy can be thought of as a (complex) access path that matches the selection condition $(day < 8/9/02 \vee rname='Joe')$.

Most current systems do not handle selection conditions with disjunction efficiently and concentrate on optimizing selections without disjunction.

14.3 THE PROJECTION OPERATION

Consider the query shown in Figure 14.2. The optimizer translates this query into the relational algebra expression $\pi_{sid,bid}Reserves$. In general the projection operator is of the form $\pi_{attr1,attr2,\ldots,attrm}(R)$. To implement projection, we have

```
SELECT DISTINCT R.sid, R.bid
FROM   Reserves R
```

Figure 14.2 Simple Projection Query

to do the following:

1. Remove unwanted attributes (i.e., those not specified in the projection).
2. Eliminate any duplicate tuples produced.

The second step is the difficult one. There are two basic algorithms, one based on sorting and one based on hashing. In terms of the general techniques listed in Section 12.2, both algorithms are instances of partitioning. While the technique of using an index to identify a subset of useful tuples is not applicable for projection, the sorting or hashing algorithms can be applied to data entries in an index, instead of to data records, under certain conditions described in Section 14.3.4.

14.3.1 Projection Based on Sorting

The algorithm based on sorting has the following steps (at least conceptually):

1. Scan R and produce a set of tuples that contain only the desired attributes.
2. Sort this set of tuples using the combination of all its attributes as the key for sorting.
3. Scan the sorted result, comparing adjacent tuples, and discard duplicates.

If we use temporary relations at each step, the first step costs M I/Os to scan R, where M is the number of pages of R, and T I/Os to write the temporary relation, where T is the number of pages of the temporary; T is $a(M)$. (The exact value of T depends on the number of fields retained and the sizes of these fields.) The second step costs *O(TlogT)* (which is also *0(MlogA1)*, of course). The final step costs T. The total cost is $O(MlogM)$. The first and third steps are straightforward and relatively inexpensive. (As noted in the chapter on sorting, the cost of sorting grows linearly with dataset size in practice, given typical dataset sizes and main memory sizes.)

Consider the projection on Reserves shown in Figure 14.2. We can scall Reserves at a cost of 1000 I/Os. If we assume that each tuple in the temporary relation created in the first step is 10 bytes long, the cost of writing this temporary relation is 250 I/Os. Suppose we have 20 buffer pages. We can sort the temporary relation in two passes at a cost of $2 \cdot 2 \cdot 250 = 1000$ I/Os. The scan required in the third step costs an additional 250 I/Os. The total cost is 2500 I/Os.

This approach can be improved on by modifying the sorting algorithm to do projection with duplicate elimination. Recall the structure of the external sorting algorithm presented in Chapter 13. The very first pass (Pass 0) involves a scan of the records that are to be sorted to produce the initial set of (internally) sorted runs. Subsequently, one or more passes merge runs. Two important modifications to the sorting algorithm adapt it for projection:

- 'We can project out unwanted attributes during the first pass (Pass 0) of sorting. If B buffer pages are available, we can read in B pages of Rand write out *(T/!vI)* . B *internally sorted* pages of the temporary relation. In fact, with a more aggressive implementation, we can write out approximately 2 . *B* internally sorted pages of the temporary relation on average. (The idea is similar to the refinement of external sorting discussed in Section 13.3.1.)
- We can eliminate duplicates during the merging passes. In fact, this modification reduces the cost of the merging passes since fewer tuples are written out in each pass. (Most of the duplicates are eliminated in the very first merging pass.)

Let us consider our example again. In the first pass we scan Reserves, at a cost of 1000 I/Os and write out 250 pages. With 20 buffer pages, the 250 pages are written out as seven internally sorted runs, each (except the last) about 40 pages long. In the second pass we read the runs, at a cost of 250 I/Os, and merge them. The total cost is 1,500 I/Os, which is much lower than the cost of the first approach used to implement projection.

14.3.2 Projection Based on Hashing

If we have a fairly large number (say, B) of buffer pages relative to the number of pages of *R,* a hash-based approach is worth considering. There are two phases: partitioning and duplicate elimination.

In the *partitioning* phase, we have one *input* buffer page and $B-1$ *output* buffer pages. The relation *R* is read into the input buffer page, one page at a time. The input page is processed as follows: For each tuple, we project out the unwanted attributes and then apply a hash function *h* to the combination of all remaining.attributes. The function *h* is chosen so that tuples are distributed uniformly to one of $B-1$ partitions; there is one output page per partition. After the projection the tuple is written to the output buffer page that it is hashed to by *h.*

At the end of the partitioning phase, we have $B-1$ partitions, each of which contains a collection of tuples that share a common hash value (computed by

applying h to all fields), and have only the desired fields. The partitioning phase is illustrated in Figure 14.3.

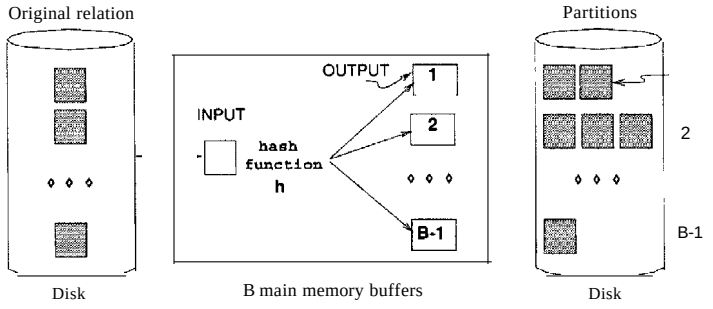


Figure 14.3 Partitioning Phase of Hash-Based Projection

Two tuples that belong to different partitions are guaranteed not to be duplicates because they have different hash values. Thus, if two tuples are duplicates, they are in the same partition. In the *duplicate elimination* phase, we read in the $B-1$ partitions one at a time to eliminate duplicates. The basic idea is to build an in-memory hash table as we process tuples in order to detect duplicates.

For each partition produced in the first phase:

1. Read in the partition one page at a time. Hash each tuple by applying hash function *h2* ($\neq h$) to the combination of all fields and then insert it into an in-memory hash table. If a new tuple hashes to the same value as some existing tuple, compare the two to check whether the new tuple is a duplicate. Discard duplicates as they are detected.

2. After the entire partition has been read in, write the tuples in the hash table (which is free of duplicates) to the result file. Then clear the in-memory hash table to prepare for the next partition.

Note that *h2* is intended to distribute the tuples in a partition across many buckets to minimize *collisions* (two tuples having the same *h2* values). Since all tuples in a given partition have the same *h* value, *h2* cannot be the same as *h!*

This hash-based projection strategy will not work well if the size of the hash table for a partition (produced in the partitioning phase) is greater than the number of available buffer pages *B*. One way to handle this *partition overflow* problem is to recursively apply the hash-based projection technique to eliminate the duplicates in each partition that overflows. That is, we divide

an overflowing partition into subpartitions, then read each subpartition into memory to eliminate duplicates.

If we assume that h distributes the tuples with perfect uniformity and that the number of pages of tuples *after* the projection (but before duplicate elimination) is T, each partition contains $\frac{T}{B-1}$ pages. (Note that the number of partitions is $B-1$ because one of the buffer pages is used to read in the relation during the partitioning phase.) The size of a partition is therefore $\frac{T}{B-1}$, and the size of a hash table for a partition is $\frac{T}{B-1} \cdot f$; where f is a *fudge factor* used to capture the (small) increase in size between the partition and a hash table for the partition. The number of buffer pages B must be greater than the partition size $\frac{T}{B-1} \cdot f$ to avoid partition overflow. This observation implies that we require approximately $B > \sqrt{f \cdot T}$ buffer pages.

Now let us consider the cost of hash-based projection. In the partitioning phase, we read R, at a cost of M I/Os. We also write out the projected tuples, a total of T pages, where T is some fraction of M, depending on the fields that are projected out. The cost of this phase is therefore $M + T$ l/Os; the cost of hashing is a CPU cost, and we do not take it into account. In the duplicate elimination phase, we have to read in every partition. The total number of pages in all partitions is T. We also write out the in-memory hash table for each partition after duplicate elimination; this hash table is part of the result of the projection, and we ignore the cost of writing out result tuples, as usual. Thus, the total cost of both phases is $M + 2T$. In our projection on Reserves (Figure 14.2), this cost is $1000 + 2 \cdot 250 = 1500$ l/Os.

14.3.3 Sorting Versus Hashing for Projections

The sorting-based approach is superior to hashing if we have many duplicates or if the distribution of (hash) values is very nonuniform. In this case, some partitions could be much larger than average, and a hash table for such a partition would not fit in memory during the duplicate elimination phase. Also, a useful side effect of using the sorting-based approach is that the result is sorted. Further, since external sorting is required for a variety of reasons, most database systems have a sorting utility, which can be used to implement projection relatively easily. For these reasons, sorting is the standard approach for projection. And perhaps due to a simplistic use of the sorting utility, unwanted attribute removal and duplicate elimination are separate steps in many systems (i.e., the basic sorting algorithm is often used without the refinements we outlined).

We observe that, if we have $B > \sqrt{T}$ buffer pages, where T is the size of the projected relation before duplicate elimination, both approaches have the

Projection in Commercial Systems: Informix uses hashing. IBMDB2, Oracle 8, and Sybase ASE use sorting. Microsoft SQL Server and Sybase ASIQ implement both hash-based and sort-based algorithms.

same I/O cost. Sorting takes two passes. In the first pass, we read M pages of the original relation and write out T pages. **In** the second pass, we read the T pages and output the result of the projection. Using hashing, in the partitioning phase, we read M pages and write T pages' worth of partitions. In the second phase, we read T pages and output the result of the projection. Thus, considerations such as CPU costs, desirability of sorted order in the result, and skew in the distribution of values drive the choice of projection method.

14.3.4 Use of Indexes for Projections

Neither the hashing nor the sorting approach utilizes any existing indexes. An existing index is useful if the key includes all the attributes we wish to retain in the projection. In this *case,* we can simply retrieve the key values from the index-without ever accessing the actual relation-and apply our projection techniques to this (much smaller) set of pages. This technique, called an *index-only scan,* and was discussed in Sections 8.5.2 and 12.3.2. **If** we have an ordered (i.e., a tree) index whose search key includes the wanted attributes as a *prefix,* we can do even better: Just retrieve the data entries in order, discarding unwanted fields, and compare adjacent entries to check for duplicates. The index-only scan technique is discussed further in Section 15.4.1.

14.4 THE JOIN OPERATION

Consider the following query:

```
SELECT *
FROM   Reserves R, Sailors S
WHERE  R.sid = S.sid
```

This query can be expressed in relational algebra using the join operation: $R \bowtie S$. The *join* operation, one of the most useful operations in relational algebra, is the primary means of combining information from two or more relations.

Joins in Commercial Systems: Sybase ASE supports index nested loop and sort-merge join. Sybase ASIQ supports page-oriented nested loop, index nested loop, simple hash, and sort-merge join, in addition to join indexes (which we discuss in Chapter 25). Oracle 8 supports page-oriented nested loops join, sort-merge join, and a variant of hybrid hash join. IBM DB2 supports block nested loop, sort-merge, and hybrid hash join. Microsoft SQL Server supports block nested loops, index nested loops, 80rt-merge, hash join, and a technique called *hash teams.* Informix supports block nested loops, index nested loops, and hybrid hash join.

Although a join can be defined as a cross-product followed by selections and projections, joins arise much more frequently in practice than plain cross-products. Further, the result of a cross-product is typically much larger than the result of a join, so it is very important to recognize joins and implement them without materializing the underlying cross-product. Joins have therefore received a lot of attention.

We now consider several alternative techniques for implementing joins. We begin by discussing two algorithms (simple nested loops and block nested loops) that essentially enumerate all tuples in the cross-product and discard tuples that do not meet the join conditions. These algorithms are instances of the simple iteration technique mentioned in Section 12.2.

The remaining join algorithms avoid enumerating the cross-product. They are instances of the indexing and partitioning techniques mentioned in Section 12.2. Intuitively, if the join condition consists of equalities, tuples in the two relations can be thought of as belonging to *partitions,* such that only tuples in the same partition can join with each other; the tuples in a partition contain the same values in the join columns. Index nested loops join scans one of the relations and, for each tuple in it, uses an index on the (join columns of the) second relation to locate tuples in the same partition. Thus, only a subset of the second relation is compared with a given tuple of the first relation, and the entire cross-product is not enumerated. The last two algorithms (sort-merge join and hash join) also take advantage of join conditions to partition tuples in the relations to be joined and compare only tuples in the same partition while computing the join, but they do not rely on a pre-existing index. Instead, they either sort or hash the relations to be joined to achieve the partitioning.

We discuss the join of two relations R and S, with the join condition $R_i = Sj$, using positional notation. (If we have more complex join conditions, the basic idea behind each algorithm remains essentially the same. \Ve discuss the details in Section 14.4.4.) We assmne M pages in R with p_R tuples per page and N

pages in S with PS tuples per page. We use R and S in our presentation of the algorithms, and the Reserves and Sailors relations for specific examples.

14.4.1 Nested Loops Join

The simplest join algorithm is a tuple-at-a-time nested loops evaluation. We scan the *outer* relation R, and for each tuple $r \in R$, we scan the entire *inner* relation S. The cost of scanning R is M I/Os. We scan S a total of $PR \cdot M$ times, and each scan costs N I/Os. Thus, the total cost is $M + PR \cdot M \cdot N$.

```
foreach tuple r ∈ R do
    foreach tuple s ∈ S do
        if ri==Sj then add (r, s) to result
```

Figure 14.4 Simple Nested Loops Join

Suppose we choose R to be Reserves and S to be Sailors. The value of M is then 1,000, PR is 100, and N is 500. The cost of simple nested loops join is $1000 + 100 \cdot 1000 \cdot 500$ page I/Os (plus the cost of writing out the result; we remind the reader again that we uniformly ignore this component of the cost). The cost is staggering: $1000 + (5 \cdot 10^7)$ I/Os. Note that each I/O costs about lams on current hardware, which means that this join will take about 140 hours!

A simple refinement is to do this join *page-at-a-time:* For each page of R, we can retrieve each page of S and write out tuples (r, s) for all qualifying tuples $r \in R\text{-}page$ and $s \in S\text{-}page$. This way, the cost is M to scan R, as before. However, S is scanned only M times, and so the total cost is $M + M \cdot N$. Thus, the page-at-a-time refinement gives us an improvement of a factor of PR. In the example join of the Reserves and Sailors relations, the cost is reduced to $1000 + 1000 \cdot 500 = 501,000$ I/Os and would take about 1.4 hours. This dramatic improvement underscores the importance of page-oriented operations for minimizing disk I/O.

From these cost formulas a straightforward observation is that we should choose the outer relation R to be the smaller of the two relations ($R \bowtie B = B \bowtie R$, as long as we keep track of field names). This choice does not change the costs significantly, however. If we choose the smaller relation, Sailors, as the outer relation, the cost of the page-at-a-time algorithm is $500 + 500 \cdot 1000 = 500,500$ I/Os, which is only marginally better than the cost of page-oriented simple nested loops join with Reserves as the outer relation.

Block Nested Loops Join

The simple nested loops join algorithm does not effectively utilize buffer pages. Suppose we have enough memory to hold the smaller relation, say, R, with at least two extra buffer pages left over. We can read in the smaller relation and use one of the extra buffer pages to scan the larger relation S. For each tuple s E 5, we check R and output a tuple $(1', s)$ for qualifying tuples s (i.e., $r_i = sj$). The second extra buffer page)s used as an output buffer. Each relation is scanned just once, for a total I/O cost of $M + N$, which is optimal.

If enough memory is available, an important refinement is to build an in-memory *hash table* for the smaller relation R. The I/O cost is still $M + N$, but the CPU cost is typically much lower with the hash table refinement.

What if we have too little memory to hold the entire smaller relation? We can generalize the preceding idea by breaking the relation R into *blocks* that can fit into the available buffer pages and scanning all of 5 for each block of R. R is the *outer* relation, since it is scanned only once, and S is the *inner* relation, since it is scanned multiple times. If we have B buffer pages, we can read in $B-2$ pages of the outer relation R and scan the inner relation S using one of the two remaining pages. We can write out tuples $(1', s)$, where r E *R-block*, s E *S-page*, and $r_i = sj$, using the last buffer page for output.

An efficient way to find **matching pairs** of tuples (i.e., tuples satisfying the join condition r_i = sj) is to build a main-memory hash table for the block of R. Because a hash table for a set of tuples takes a little more space than just the tuples themselves, building a hash table involves a trade-off: The effective block size of R, in terms of the number of tuples per block, is reduced. Building a hash table is well worth the effort. The block nested loops algorithm is described in Figure 14.5. Buffer usage in this algorithm is illustrated in Figl.lre 14.6.

```
foreach block of B-2 pages of R do
    foreach page of 5 do {
        for all matching in-memory tuples r E R-block and s E S-page,
        add (1', s) to result
    }
```

Figure 14.5 Block Nested Loops Join

The cost of this strategy is M 1/0s for reading in R (which is scanned only once). 5 is scanned a total of $\lceil \frac{M}{B-2} \rceil$ times-ignoring the extra space required per page due to the in-memory hash table--·and each scan costs N l/Os. The total cost is thus $M + N \cdot \lceil \frac{M}{B-2} \rceil$.

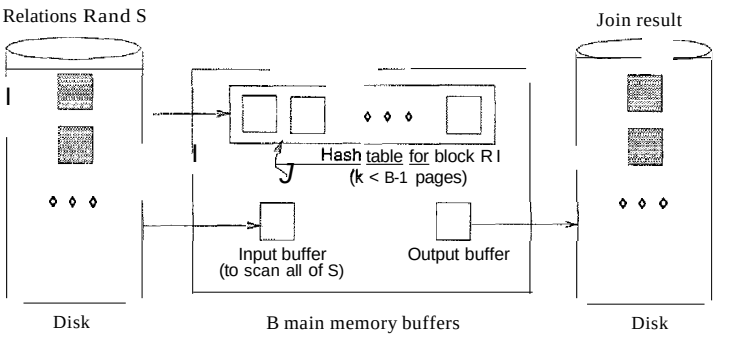


Figure 14.6 Buffer Usage in Block Nested Loops Join

Consider the join of the Reserves and Sailors relations. Let us choose Reserves to be the outer relation R and assume we have enough buffers to hold an in-memory hash table for 100 pages of Reserves (with at least two additional buffers, of course). We have to scan Reserves, at a cost of 1000 I/Os. For each 100-page block of Reserves, we have to scan Sailors. Therefore, we perform 10 scans of Sailors, each costing 500 I/Os. The total cost is $1000 + 10 \cdot 500 = 6000$ I/Os. If we had only enough buffers to hold 90 pages of Reserves, we would have to scan Sailors $\lceil 1000/90 \rceil = 12$ times, and the total cost would be $1000 + 12 \cdot 500 = 7000$ I/Os.

Suppose we choose Sailors to be the outer relation R instead. Scanning Sailors costs 500 I/Os. We would scan Reserves $\lceil 500/100 \rceil = 5$ times. The total cost is $500 + 5 \cdot 1{,}000 = 5500$ I/Os. If instead we have only enough buffers for 90 pages of Sailors, we would scan Reserves a total of $\lceil 500/90 \rceil = 6$ times. The total cost in this case is $500 + 6 \cdot 1000 = 6500$ I/Os. We note that the block nested loops join algorithm takes a little over a minute on our running example, assuming 10ms per I/O as before.

Impact of Blocked Access

If we consider the effect of blocked access to several pages, there is a fundamental change in the way we allocate buffers for block nested loops. Rather than using just one buffer page for the inner relation, the best approach is to split the buffer pool evenly between the two relations. This allocation results in more passes over the inner relation, leading to more page fetches. However, the time spent on *seeking* for pages is dramatically reduced.

The technique of double buffering (discussed in Chapter 13 in the context of sorting) can also be used, but we do not discuss it further.

Index Nested Loops Join

If there is an index on one of the relations on the join attribute(s), we can take advantage of the index by making the indexed relation be the inner relation. Suppose we have a suitable index on *S;* Figure 14.7 describes the index nested loops join algorithm.

```
foreach tuple r E R do
    foreach tuple s E S where r_i == s_j
        add ⟨r, s) to result
```

Figure 14.7 Index Nested Loops Join

For each tuple $r \in R$, we use the index to retrieve matching tuples of S. Intuitively, we compare r only with tuples of S that are in the same *partition,* in that they have the same value in the join column. Unlike the other nested loops join algorithms, therefore, the index nested loops join algorithm does not enumerate the cross-product of Rand S. The cost of scanning R is M, as before. The cost of retrieving matching S tuples depends on the kind of index and the number of matching tuples; for each R tuple, the cost is as follows:

1. If the index on S is a B+ tree index, the cost to find the appropriate leaf is typically 2---4 l/Os. If the index is a hash index, the cost to find the appropriate bucket is 1-2 1/Os.
2. Once we find the appropriate leaf or bucket, the cost of retrieving matching S tuples depends on whether the index is clustered. If it is, the cost per outer tuple $r \in R$ is typically just one more I/O. If it is not clustered, the cost could be one I/O per matching S-tuple (since each of these could be on a different page in the worst case).

As an example, suppose that we have a hash-based index using Alternative (2) on the *sid* attribute of Sailors and that it takes about 1.2 l/Os on average[2] to retrieve the appropriate page of the index. Since *sid* is a key for Sailors, we have at most one matching tuple. Indeed, *sid* in Reserves is a foreign key referring to Sailors, and therefore we have *exactly* one matching Sailors tuple for each Reserves tuple. Let us consider the cost of scanning Reserves and using the index to retrieve the matching Sailors tuple for each Reserves tuple. The cost of scanning Reserves is 1000. There are 100 . 1000 tuples in Reserves. For each of these tuples, retrieving the index page containing the rid of the matching Sailors tuple costs 1.2 l/Os (on average); in addition, we have to retrieve the Sailors page containing the qualifying tuple. Therefore, we have

2This is a typical cost for hash-based indexes,

100,000 ·(1 + 1.2) l/Os to retrieve matching Sailors tuples. The total cost is 221,000 l/Os.

As another example, suppose that we have a hash-based index using Alternative (2) on the *sid* attribute of Reserves. Now we can scan Sailors (500 l/Os), and for each tuple, use the index to retrieve matching Reserves tuples. We have a total of 80 . 500 Sailors tuples, and each tuple could match with either zero or more Reserves tuples; a sailor may have no reservations or several. For each Sailors tuple, we can retrieve the index page containing the rids of matching Reserves tuples (assuming that we have at most one such index page, which is a reasonable guess) in 1.2 l/Os on average. The total cost thus far is 500 + 40,000 . 1.2 = 48,500 l/Os.

In addition, we have the cost of retrieving matching Reserves tuples. Since we have 100,000 reservations for 40,000 Sailors, assuming a uniform distribution we can estimate that each Sailors tuple matches with 2.5 Reserves tuples on average. If the index on Reserves is clustered, and these matching tuples are typically on the same page of Reserves for a given sailor, the cost of retrieving them is just one I/O per Sailor tuple, which adds up to 40,000 extra l/Os. If the index is not clustered, each matching Reserves tuple may well be on a different page, leading to a total of 2.5 . 40,000 l/Os for retrieving qualifying tuples. Therefore, the total cost can vary from 48,500+40,000=88,500 to 48,500+100,000=148,500 l/Os. Assuming 10ms per I/O, this would take about 15 to 25 minutes.

So, even with an unclustered index, if the number of matching inner tuples for each outer tuple is small (on average), the cost of the index nested loops join algorithm is likely to be much less than the cost of a simple nested loops join.

14.4.2 Sort-Merge Join

The basic idea behind the sort-merge **join** algorithm is to *sort* both relations on the join attribute and then look for qualifying tuples $r \in R$ and $s \in S$ by essentially *merging* the two relations. The sorting step groups all tuples with the same value in the join column and thus makes it easy to identify partitions, or groups of tuples with the same value, in the join column. We exploit this partitioning by comparing the R tuples in a partition with only the S tuples in the same partition (rather than with all S tuples), thereby avoiding enumeration of the cross-product of R and S. (This partition-based approach works only for equality join conditions.)

The external sorting algorithm discussed in Chapter 13 can be used to do the sorting, and of course, if a relation is already sorted on the join attribute, we

need not sort it again. We now consider the merging step in detail: We scan the relations *Rand S)* looking for qualifying tuples (i.e., tuples *Tr* in *Rand Ts* in *S* such that $Tr_i = Tsj$). The two scans start at the first tuple in each relation. We advance the scan of *R* as long as the current *R* tuple is less than the current *S* tuple (with respect to the values in the join attribute). Similarly, we advance the scan of *S* as long as the current *S* tuple is less than the current *R* tuple. We alternate between such advances until we find an *R* tuple *Tr* and a *S* tuple *Ts* with $Tr_i = TSj'$

When we find tuples *Tr* and *Ts* such that $Tri = Tsj$, we need to output the joined tuple. In fact, we could have several *R* tuples and several *S* tuples with the same value in the join attributes as the current tuples *Tr* and *Ts*. We refer to these tuples as the *current R partition* and the *current S partition.* For each tuple *r* in the current *R* partition, we scan all tuples *s* in the current *S* partition and output the joined tuple *(r, s).* We then resume scanning Rand *S,* beginning with the first tuples that follow the partitions of tuples that we just processed.

The sort-merge join algorithm is shown in Figure 14.8. We assign only tuple values to the variables *Tr, Ts,* and *Gs* and use the special value *eof* to denote that there are no more tuples in the relation being scanned. Subscripts identify fields, for example, Tr_i denotes the ith field of tuple *Tr*. If *Tr* has the value *eof,* any comparison involving *Tri* is defined to evaluate to false.

We illustrate sort-merge join on the Sailors and Reserves instances shown in Figures 14.9 and 14.10, with the join condition being equality on the *sid* attributes.

These two relations are already sorted on *sid,* and the merging phase of the sort-merge join algorithm begins with the scans positioned at the first tuple of each relation instance. We advance the scan of Sailors, since its *sid* value, now 22, is less than the *sid* value of Reserves, which is now 28. The second Sailors tuple has *sid* = 28, which is equal to the *sid* value of the current Reserves tuple. Therefore, we now output a result tuple for each pair of tuples, one from Sailors and one from Reserves, in the current partition (i.e., with *sid* = 28). Since we have just one Sailors tuple with *sid* = 28 and two such Reserves tuples, we write two result tuples. After this step, we position the scan of Sailors at the first tuple after the partition with *sid* = 28, which has *sid* = 31. Similarly, we position the scan of Reserves at the first tuple with *sid* = 31. Since these two tuples have the same *sid* values, we have found the next matching partition, and we must write out the result tuples generated from this partition (there are three such tuples). After this, the Sailors scan is positioned at the tuple with *sid* = 36, and the Reserves scan is positioned at the tuple with *sid* = 58. The rest of the merge phase proceeds similarly.

```
proc smjoin(R, B, 'Ri = Sj')

if R not sorted on attribute i, sort it;
if B not sorted on attribute j, sort it;

Tr = first tuple in R;                                  // ranges over R
Ts = first tuple in B;                                  // ranges over S
Gs = first tuple in S;                          // start of current S-partition

while Tr ≠ eof and Gs ≠ eof do {

    while Tri < GSj do
        Tr = next tuple in Rafter Tr;                   // continue scan of R

    while Tri > GSj do
        Gs = next tuple in S after Gs                   // continue scan of B

    Ts = Gs;                                    // Needed in case Tri ≠ GSj
    while Tri == GSj do {                       // process current R partition
        Ts = Gs;                                        // reset S partition scan
        while TSj == Tri do {                           // process current R tuple
            add (Tr, Ts) to result;                     // output joined tuples
            Ts = next tuple in S after Ts;} // advance S partition scan
        Tr = next tuple in Rafter Tr;                   // advance scan of R
        }                                       // done with current R partition

    Gs = Ts;                                // initialize search for next S partition

    }
```

Figure 14.8 Sort-Merge Join

sid	sname	rating	age
22	dustin	7	45.0
28	yuppy	9	35.0
31	lubber	8	55.5
36	lubber	6	36.0
44	guppy	5	35.0
58	rusty	10	35.0

Figure 14.9 An Instance of Sailors

sid	bid	day	rname
28	103	12/04/96	guppy
28	103	11/03/96	}'uppy
31	101	10/10/96	dustin
31	102	10/12/96	lubber
31	101	10/11/96	lubber
58	103	11/12/96	dustin

Figure 14.10 An Instance of Reserves

In general, we have to scan a partition of tuples in the second relation as often as the number of tuples in the corresponding partition in the first relation. The first relation in the example, Sailors, has just one tuple in each partition. (This is not happenstance but a consequence of the fact that *sid* is a key—this example is a key-foreign key join.) In contrast, suppose that the join condition is changed to be *sname=7'name.* Now, both relations contain more than one tuple in the partition with *sname=mame='lubber'.* The tuples with *rname='lubber'* in Reserves have to be scanned for each Sailors tuple with *sname='lubber'.*

Cost of Sort-Merge Join

The cost of sorting R is $O(MlogM)$ ancl the cost of sorting S is $O(NlogN)$. The cost of the merging phase is $M + N$ if no S partition is scanned multiple times (or the necessary pages are found in the buffer after the first pass). This approach is especially attractive if at least one relation is already sorted on the join attribute or has a clustered index on the join attribute.

Consider the join of the relations Reserves and Sailors. Assuming that we have 100 buffer pages (roughly the same number that we assumed were available in our discussion of block nested loops join), we can sort Reserves in just two passes. The first pass produces 10 internally sorted runs of 100 pages each. The second pass merges these 10 runs to produce the sorted relation. Because we read and write Reserves in each pass, the sorting cost is $2 \cdot 2 \cdot 1000 = 4000$ l/Os. Similarly, we can sort Sailors in two passes, at a cost of $2 \cdot 2 \cdot 500 = 2000$ l/Os. In addition, the seconcl phase of the sort-merge join algorithm requires an additional scan of both relations. Thus the total cost is $4000 + 2000 + 1000 + 500 = 7500$ l/Os, which is similar to the cost of the block nested loops algorithm.

Suppose that we have only 35 buffer pages. We can still sort both Reserves and Sailors in two passes, and the cost of the sort-merge join algorithm remains at 7500 l/Os. However, the cost of the block nested loops join algorithm is more than 15,000 l/Os. On the other hand, if we have 300 buffer pages, the cost of the sort-merge join remains at 7500 I/Os, whereas the cost of the block nested loops join drops to 2500 l/Os. (We leave it to the reader to verify these numbers.)

We note that multiple scans of a partition of the second relation are potentially expensive. In our example, if the number of Reserves tuples in a repeatedly scanned partition is small (say, just a few pages), the likelihood of finding the entire partitiOli in the buffer pool on repeated scans is very high, and the I/O cost remains essentially the same as for a single scan. However, if many pages

of Reserves tuples are in a given partition, the first page of such a partition may no longer be in the buffer pool when we request it a second time (after first scanning all pages in the partition; remember that each page is unpinned as the scan moves past it). In this case, the I/O cost could be as high as the number of pages in the Reserves partition times the number of tuples in the corresponding Sailors partition!

In the worst-case scenario, the merging phase could require us to read the complete second relation for each *tuple* in the first relation, and the number of I/Os is $O(M \cdot N)$ I/Os! (This scenario occurs when all tuples in both relations contain the same value in the join attribute; it is extremely unlikely.)

In practice, the I/O cost of the merge phase is typically just a single scan of each relation. A single scan can be guaranteed if at least one of the relations involved has no duplicates in the join attribute; this is the case, fortunately, for key–foreign key joins, which are very common.

A Refinement

We assumed that the two relations are sorted first and then merged in a distinct pass. It is possible to improve the sort-merge join algorithm by combining the merging phase of sorting with the merging phase of the join. First, we produce sorted runs of size B for both Rand 5. If $B > \sqrt{L}$, where L is the size of the larger relation, the number of runs per relation is less than $\sqrt{L}$. Suppose that the number of buffers available for the merging phase is at least $2\sqrt{L}$; that is, more than the total number of runs for Rand 5. We allocate one buffer page for each run of R *and* one for each run of 5. We then merge the runs of R (to generate the sorted version of R), merge the runs of 5, and merge the resulting Rand 5 streams as they are generated; we apply the join condition as we merge the Rand S streams and discard tuples in the cross--product that do not meet the join condition.

Unfortunately, this idea increases the number of buffers required to $2\sqrt{L}$. However, by using the technique discussed in Section 13.3.1 we can produce sorted runs of size approximately $2 \cdot B$ for both Rand 5. Consequently, we have fewer than $\sqrt{L}/2$ runs of each relation, given the assumption that $B > \sqrt{L}$. Thus, the total number of runs is less than $\sqrt{L}$, that is, less than B, and we can combine the merging phases with no need for additional buffers.

This approach allows us to perform a sort-merge join at the cost of reading and writing Rand S in the first pass and reading Rand 5 in the second pass. The total cost is thus $3 \cdot (M + N)$. In our example, the cost goes down from 7500 to 4500 I/Os.

Blocked Access and Double-Buffering

The blocked I/O and double-buffering optimizations, discussed in Chapter 13 in the context of sorting, can be used to speed up the merging pass as well as the sorting of the relations to be joined; we do not discuss these refinements.

14.4.3 Hash Join

The hash join algorithm, like the sort-merge join algorithm, identifies partitions in Rand *S* in a partitioning phase and, in a subsequent probing phase, compares tuples in an *R* partition only with tuples in the corresponding *5* partition for testing equality join conditions. Unlike sort-merge join, hash join uses hashing to identify partitions rather than sorting. The partitioning (also called building) phase of hash join is similar to the partitioning in hash-based projection and is illustrated in Figure 14.3. The probing (sometimes called matching) phase is illustrated in Figure 14.11.

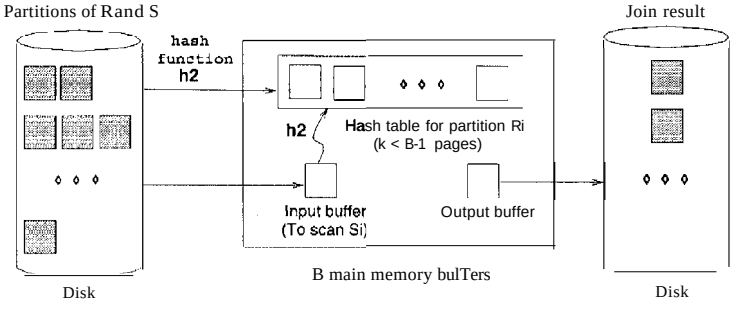


Figure 14.11 Probing Phase of Hash Join

The idea is to hash *both* relations on the join attribute, using the *same* hash function *h.* If we hash each relation (ideally uniformly) into *k* partitions, we are assured that *R* tuples in partition i can join only with *S* tuples in the same partition i. This observation can be used to good effect: We can read in a (complete) partition of the smaller relation *R* and scan just the corresponding partition of *S* for matches. We never need to consider these Rand *S* tuples again. Thus, once Rand *S* are partitioned, we can perform the join by reading in *R* and *5* just once, provided enough memory is available to hold all the tuples in any given partition of *R.*

In practice we build an in-memory hash table for the *R* partition, using a ha'3h function *h2* that is different from *h* (since *h2* is intended to distribute tuples in a partition based on *h),* to reduce CPU costs. We need enough memory to hold this hash table, which is a little larger than the *R* partition itself.

The hash join algorithm is presented in Figure 14.12. (There are several variants on this idea; this version is called *Grace hash join* in the literature.) Consider the cost of the hash join algorithm. In the partitioning phase, we have to scan both R and S once and write them out once. The cost of this phase is therefore 2(M + N). In the second phase, we scan each partition once, assuming no partition overflows, at a cost of M + N I/Os. The total cost is therefore 3(M + N), given our assumption that each partition fits into memory in the second phase. On our example join of Reserves and Sailors, the total cost is 3 · (500 + 1000) = 4500 I/Os, and assuming 10ms per I/O, hash join takes under a minute. Compare this with simple nested loops join, which took about 140 hours--this difference underscores the importance of using a good join algorithm.

```
// Partition R into k partitions
foreach tuple r ∈ R do
    read r and add it to buffer page h(ri);        // flushed as page fills

// Partition S into k partitions
foreach tuple s ∈ S do
    read s and add it to buffer page h(sj);        // flushed as page fills

// Probing phase
for l = 1, ..., k do {

    // Build in-memory hash table for Rz, using h2
    foreach tuple r ∈ partition Rz do
        read r and insert into hash table using h2(ri) ;

    // Scan Sz and probe for matching Rz tuples
    foreach tuple s ∈ partition Sz do {
        read s and probe table using h2(sj);
        for matching R tuples r, output ⟨r, s⟩ };

    clear hash table to prepare for next partition;
    }
```

Figure 14.12 Hash Join

Memory Requirements and Overflow Handling

To increase the chances of a given partition fitting into available memory in the probing phase, we must minimize the size of a partition by maximizing the number of partitions. In the partitioning phase, to partition R (similarly,

8) into k partitions, we need at least k output buffers and one input buffer. Therefore, given B buffer pages, the maximum number of partitions is $k = B - 1$. Assuming that partitions are equal in size, this means that the size of each R partition is $\frac{M}{B-1}$ (as usual, M is the number of pages of R). The number of pages in the (in-memory) hash table built during the probing phase for a partition is thus $\frac{f \cdot M}{B-1}$, where f is a *fudge factor* used to capture the (small) increase in size between the partition and a hash table for the partition.

During the probing phase, in addition to the hash table for the R partition, we require a buffer page for scanning the 8 partition and an output buffer. Therefore, we require $B > \frac{f \cdot M}{B-1} + 2$. *We need approximately* $B > \sqrt{f} \cdot M$ for the hash join algorithm to perform well.

Since the partitions of R are likely to be close in size but not identical, the largest partition is somewhat larger than $\frac{M}{B-1}$, and the number of buffer pages required is a little more than $B > \sqrt{f} \cdot M$. There is also the risk that, if the hash function h does not partition R uniformly, the hash table for one or more R partitions may not fit in memory during the probing phase. This situation can significantly degrade performance.

As we observed in the context of hash-based projection, one way to handle this *partition overflow* problem is to recursively apply the hash join technique to the join of the overflowing R partition with the corresponding 8 partition. That is, we first divide the Rand 8 partitions into subpartitions. Then, we join the subpartitions pairwise. All subpartitions of R probably fit into memory; if not, we apply the hash join technique recursively.

Utilizing Extra Memory: Hybrid Hash Join

The minimum amount of memory required for hash join is $B > \sqrt{f} \cdot M$. If more memory is available, a variant of hash join called **hybrid hash join** offers better performance. Suppose that $B > f \cdot (lYI/k)$, for some integer k. This means that, if we divide R into k partitions of size M/k, an in-memory hash table can be built for each partition. To partition R (similarly, 5) into k partitions, we need k output buHers and one input buffer; that is, $k + 1$ pages. This leaves us with $B - (k + 1)$ extra pages during the partitioning phase.

Suppose that $B - (k + 1) > f \cdot (M/k)$. That is, we have enough extra memory during the partitioning phase to hold an in-memory hash table for a partition of R. The idea behind hybrid hash join is to build an in-memory hash table for the first partition of R during the partitioning phase, which means that we do not write this partition to disk. Similarly, while partitioning 8, rather than write out the tuples in the first partition of 5, we can directly probe the

in-memory table for the first R partition and write out the results. At the end of the partitioning phase, we have completed the join of the first partitions of R and S, in addition to partitioning the two relations; in the probing phase, we join the remaining partitions as in hash join.

The savings realized through hybrid hash join is that we avoid writing the first partitions of R and S to disk during the partitioning phase and reading them in again during the probing phase. Consider our example, with 500 pages in the smaller relation R and 1000 pages in S.[3] If we have $B = 300$ pages, we can easily build an in-memory hash table for the first R partition while partitioning R into two partitions. During the partitioning phase of R, we scan R and write out one partition; the cost is 500 + 250 if we assume that the partitions are of equal size. We then scan S and write out one partition; the cost is 1000 + 500. In the probing phase, we scan the second partition of R and of S; the cost is 250 + 500. The total cost is 750 + 1500 + 750 = 3000. In contrast, the cost of hash join is 4500.

If we have enough memory to hold an in-memory hash table for all of R, the savings are even greater. For example, if $B > f . N + 2$, that is, $k = 1$, we can build an in-memory hash table for all of R. This llleans that we read R only once, to build this hash table, and read S once, to probe the R hash table. The cost is 500 + 1000 = 1500.

Hash Join Versus Block Nested Loops Join

While presenting the block nested loops join algorithm, we briefly discussed the idea of building an in-memory hash table for the inner relation. We now compare this (more CPU-efficient) version of block nested loops join with hybrid hash join.

If a hash table for the entire smaller relation fits in memory, the two algorithms are identical. If both relations are large relative to the available buffer size, we require several passes over one of the relations in block nested loops join; hash join is a more effective application of hashing techniques in this case. The I/O saved in this case by using the hash join algorithm in comparison to a block nested loops join is illustrated in Figure 14.13. In the latter, we read in all of S for each block of R; the I/O cost corresponds to the whole rectangle. In the hash join algorithm, for each block of R, we read only the corresponding block of S; the I/O cost corresponds to the shaded areas in the figure. This difference in I/O due to scans of S is highlighted in the figure.

3It is unfortunate, that in our running example, the smaller relation, which we denoted by the variable R in our discussion of hash join, is in fact the Sailors relation, which is more naturally denoted by S!

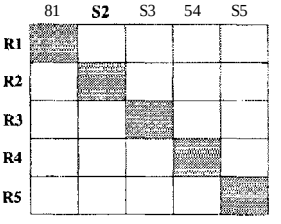


Figure 14.13 Hash Join Vs. Block Nested Loops for Large Relations

We note that this picture is rather simplistic. It does not capture the costs of scanning R in the block nested loops join and the partitioning phase in the hash join, and it focuses on the cost of the probing phase.

Hash Join Versus Sort-Merge Join

Let us compare hash join with sort-merge join. If we have $B > \sqrt{M}$ buffer pages, where M is the number of pages in the *smaller* relation and we assume uniform partitioning, the cost of hash join is $3(M + N)$ l/Os. If we have $B > \sqrt{N}$ buffer pages, where N is the number of pages in the *larger* relation, the cost of sort-merge join is also $3(M + N)$, as discussed in Section 14.4.2. A choice between these techniques is therefore governed by other factors, notably:

- If the partitions in hash join are not uniformly sized, hash join could cost more. Sort-merge join is less sensitive to such data skew.
- If the available number of buffers falls between $\sqrt{M}$ and $\sqrt{N}$, hash join costs less than sort-merge join, since we need only enough memory to hold partitions of the smaller relation, whereas in sort-merge join the memory requirements depend on the size of the larger relation. The larger the difference in size between the two relations, the more important this factor becomes.
- Additional considerations include the fact that the result is sorted in sort-merge join.

14.4.4 General Join Conditions

We have discussed several join algorithms for the case of a simple equality join condition. Other important cases include a join condition that involves equalities over several attributes and inequality conditions. To illustrate the case of several equalities, we consider the join of Reserves R and Sailors S with the join condition *R.sid=S.s'id 1\ R.rname=S.sname:*

- For index nested loops join, we can **build** an index on Reserves on the combination of fields *(R.sid, R.rname)* and treat Reserves as the inner relation. We can also use an existing index on this combination of fields, or on *R.sid*, or on *R.marne*. (Similar remarks hold for the choice of Sailors as the inner relation, of course.)
- For sort-merge join, we sort Reserves on the combination of fields ⟨*sid, rname*⟩ and Sailors on the combination of fields ⟨*sid, snarne)*. Similarly, for hash join, we partition on these combinations of fields.
- The other join algorithms we discussed are essentially unaffected.

If we have an inequality comparison, for example, a join of Reserves R and Sailors 5 with the join condition *R.rnarne* < *S.sname:*

- We require a B+ tree index for index nested loops join.
- Hash join and sort-merge join are not applicable.
- The other join algorithms we discussed are essentially unaffected.

Of course, regardless of the algorithm, the number of qualifying tuples in an inequality join is likely to be much higher than in an equality join.

We conclude our presentation of joins with the observation that no one join algorithm is uniformly superior to the others. The choice of a good algorithm depends on the sizes of the relations being joined, available access methods, and the size of the buffer pool. This choice can have a considerable impact on performance because the difference between a good and a bad algorithm for a given join can be enormous.

14.5 THE SET OPERATIONS

We now briefly consider the implementation of the set operations $R \cap S$, $R \times S$, $R \cup S$, and $R - S$. From an implementation standpoint, intersection and cr08S-product can be seen as special cases of join (with equality on all fields as the join condition for intersection, and with no join condition for cross-product). Therefore, we will not discuss them further.

The main point to acldress in the implementation of union is the elimination of duplicates. Set-difference can also be implemented using a variation of the techniques for duplicate elimination. (Union and difference queries on a single relation can be thought of as a selection query with a complex selection condition. The techniques discussecl in Section 14.2 are applicable for such queries.)

There are two implementation algorithms for union and set-difference, again based on sorting and hashing. Both algorithms are instances of the partitioning technique mentioned in Section 12.2.

14.5.1 Sorting for Union and Difference

To implement $R \cup S$:

1. Sort R using the combination of all fields; similarly, sort S.
2. Scan the sorted R and S in parallel and merge them, eliminating duplicates.

As a refinement, we can produce sorted runs of R and S and merge these runs in parallel. (This refinement is similar to the one discussed in detail for projection.) The implementation of $R - S$ is similar. During the merging pass, we write only tuples of R to the result, after checking that they do not appear in S.

14.5.2 Hashing for Union and Difference

To implement $R \cup S$:

1. Partition R and S using a hash function h.
2. Process each partition i as follows:
   - Build an in-memory hash table (using hash function $h2 \neq h$) for S_i.
   - Scan R_i. For each tuple, probe the hash table for S_i. If the tuple is in the hash table, discard it; otherwise, add it to the table.
   - Write out the hash table and then clear it to prepare for the next partition.

To implement $R - S$, we proceed similarly. The difference is in the processing of a partition. After building an in-memory hash table for S_i, we scan R_i. For each R_i tuple, we probe the hash table; if the tuple is not in the table, we write it to the result.

14.6 AGGREGATE OPERATIONS

The SQL query shown in Figure 14.14 involves an *aggregate operation*, AVG. The other aggregate operations supported in SQL-92 are MIN, MAX, SUM, and COUNT.

```
SELECT  AVG(S.age)
FROM    Sailors S
```

Figure 14.14 Simple Aggregation Query

The basic algorithm for aggregate operators consists of scanning the entire Sailors relation and maintaining some **running information** about the scanned tuples; the details are straightforward. The running information for each aggregate operation is shown in Figure 14.15. The cost of this operation is the cost of scanning all Sailors tuples.

Aggregate Operation	*Running Inforrniation*
SUM	*Total* of the values retrieved
AVG	*(Total, Count)* of the values retrieved
COUNT	*Count* of values retrieved.
MIN	Smallest value retrieved
MAX	Largest value retrieved

Figure 14.15 Running Information for Aggregate Operations

Aggregate operators can also be used in combination with a GROUP BY clause. If we add GROUP BY *rating* to the query in Figure 14.14, we would have to compute the average age of sailors for each *rating* group. For queries with grouping, there are two good evaluation algorithms that do not rely on an existing index: One algorithm is based on sorting and the other is based on hashing. Both algorithms are instances of the partitioning technique mentioned in Section 12.2.

The *sorting* approach is simple-we sort the relation on the grouping attribute *(rating)* and then scan it again to compute the result of the aggregate operation for each group. The second step is similar to the way we implement aggregate operations without grouping, with the only additional point being that we have to watch for group boundaries. (It is possible to refine the approach by doing aggregation as part of the sorting step; we leave this as an exercise for the reader.) The I/O cost of this approach is just the cost of the sorting algorithm.

In the *hashing* approach we build a hash table (in main memory, if possible) on the grouping attribute. The entries have the form *(gTOuping-value, running-info)*. The running information depends on the aggregate operation, as per the discussion of aggregate operations without grouping. As we scan the relation, for each tuple, we probe the hash table to find the entry for the group to which the tuple belongs and update the running information. 'When the hash table is cOlnplete, the entry for a grouping value can be used to compute the answer tuple for the corresponding group in the obvious way. If the hash table fits in

memory, which is likely because each entry is quite small and there is only one entry per grouping value, the cost of the hashing approach is $O(M)$, where M is the size of the relation.

If the relation is so large that the hash table does not fit in memory, we can partition the relation using a hash function *h* on *gTOuping-value.* Since all tuples with a given grouping value are in the same partition, we can then process each partition independently by building an in-memory hash table for the tuples in it.

14.6.1 Implementing Aggregation by Using an Index

The technique of using an index to select a subset of useful tuples is not applicable for aggregation. However, under certain conditions, we can evaluate aggregate operations efficiently by using the data entries in an index instead of the data records:

- If the search key for the index includes all the attributes needed for the aggregation query, we can apply the techniques described earlier in this section to the set of data entries in the index, rather than to the collection of data records and thereby avoid fetching data records.
- If the GROUP BY clause attribute list forms a prefix of the index search key and the index is a tree index, we can retrieve data entries (and data records, if necessary) in the order required for the grouping operation and thereby avoid a sorting step.

A given index may support one or both of these techniques; both are examples of *index-only* plans. We discuss the use of indexes for queries with grouping and aggregation in the context of queries that also include selections and projections in Section 15.4.1.

14.7 THE IMPACT OF BUFFERING

In implementations of relational operators, effective use of the buffer pool is very important, and we explicitly considered the size of the buffer pool in determining algorithm parameters for several of the algorithms discussed. There are three main points to note:

1. If several operations execute concurrently, they share the buffer pool. This effectively reduces the number of buffer pages available for each operation.
2. If tuples are accessed using an index, especially an unclustered index, the likelihood of finding a page in the buffer pool if it is requested multiple

times depends (in a rather unpredictable way, unfortunately) on the size of the buffer pool and the replacement policy. Further, if tuples are accessed using an unclustered index, each tuple retrieved is likely to require us to bring in a new page; therefore, the buffer pool fills up quickly, leading to a high level of paging activity.

3. If an operation has a *pattern* of repeated page accesses, we can increase the likelihood of finding a page in memory by a good choice of replacement policy or by *reserving* a sufficient number of buffers for the operation (if the buffer manager provides this capability). Several examples of such patterns of repeated access follow:

   - Consider a simple nested loops join. For each tuple of the outer relation, we repeatedly scan all pages in the inner relation. If we have enough buffer pages to hold the entire inner relation, the replacement policy is irrelevant. Otherwise, the replacement policy becomes critical. With LRU, we will *never* find a page when it is requested, because it is paged out. This is the *sequential flooding* problem discussed in Section 9.4.1. With MRU, we obtain the best buffer utilization—the first $B-2$ pages of the inner relation always remain in the buffer pool. (B is the number of buffer pages; we use one page for scanning the outer relation[4] and always replace the last page used for scanning the inner relation.)
   - In a block nested loops join, for each block of the outer relation, we scan the entire inner relation. However, since only one unpinned page is available for the scan of the inner relation, the replacement policy makes no difference.
   - In an index nested loops join, for each tuple of the outer relation, we use the index to find matching inner tuples. If several tuples of the outer relation have the same value in the join attribute, there is a repeated pattern of access on the inner relation; we can maximize the repetition by sorting the outer relation on the join attributes.

14.8 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Consider a simple selection query of the form $\sigma_{R.attr\ \mathbf{op}\ value}(R)$. What are the alternative access paths in each of these cases: (i) there is no index and the file is not sorted, (ii) there is no index but the file is sorted. (Section 14.1)

4Think about the sequence of pins and unpins used to achieve this.

- If a B+ tree index matches the selection condition, how does clustering affect the cost? Discuss this in terms of the selectivity of the condition. **(Section 14.1)**
- Describe *conjunctive normal form* for general selections. Define the terms *conjunct* and *disjunct*. Under what conditions does a general selection condition match an index? **(Section 14.2)**
- Describe the various implementation options for general selections. **(Section 14.2)**
- Discuss the use of sorting versus hashing to eliminate duplicates during projection. **(Section 14.3)**
- When can an index be used to implement projections, without retrieving actual data records? When does the index additionally allow us to eliminate duplicates without sorting or hashing? **(Section 14.3)**
- Consider the join of relations Rand 5. Describe *simple nested loops join* and *block nested loops join.* What are the similarities and differences? How does the latter reduce I/O costs? Discuss how you would utilize buffers in block nested loops. **(Section 14.4.1)**
- Describe *index nested loops join.* How does it differ from block nested loops join? **(Section 14.4.1)**
- Describe sort-merge join of Rand 5. What join conditions are supported? What optimizations are possible beyond sorting both Rand 5 on the join attributes and then doing a merge of the two? In particular, discuss how steps in sorting can be combined with the merge pass. **(Section 14.4.2)**
- What is the idea behind *hash join?* What is the additional optimization in *hybrid hash join?* **(Section 14.4.3)**
- Discuss how the choice of join algorithm depends on the number of buffer pages available, the sizes of Rand 5, and the indexes available. Be specific in your discussion and refer to cost formulas for the I/O cost of each algorithm. **(Sections** 14.12 Section **14.13)**
- How are general join conditions handled? **(Section 14.4.4)**
- \Vhy are the set operations $R \cap 5$ and R x S special cases of joins? What is the similarity between the set operations $Ru5$ and R - 5? **(Section 14.5)**
- Discuss the use of sorting versus hashing in implementing $Ru5$ and R - S. Compare this with the ilnplementation of projection. **(Section 14.5)**
- Discuss the use of *running information* in implementing aggregate operations. Discuss the use of sorting versus hashing for dealing with grouping. **(Section 14.6)**

- Under what conditions can we use an index to implement aggregate operations without retrieving data records? Under what conditions do indexes allow us to avoid sorting or hashing? (Section 14.6)
- Using the cost formulas for the various relational operator evaluation algorithms, discuss which operators are most sensitive to the number of available buffer pool pages. How is this number influenced by the number of operators being evaluated concurrently? (Section 14.7)
- Explain how the choice of a good buffer pool replacement policy can influence overall performance. Identify the patterns of access in typical relational operator evaluation and how they influence the choice of replacement policy. (Section 14.7)

EXERCISES

Exercise 14.1 Briefly answer the following questions:

1. Consider the three basic techniques, *iteration*, *indexing*, and *partitioning*, and the relational algebra operators *selection*, *projection*, and *join*. For each technique-operator pair, describe an algorithm based on the technique for evaluating the operator.
2. Define the term *most selective access path for a query*.
3. Describe *conjunctive normal form*, and explain why it is important in the context of relational query evaluation.
4. When does a general selection condition *match* an index? What is a *primary term* in a selection condition with respect to a given index?
5. How does hybrid hash join improve on the basic hash join algorithm?
6. Discuss the pros and cons of hash join, sort-merge join, and block nested loops join.
7. If the join condition is not equality, can you use sort-merge join? Can you use hash join? Can you use index nested loops join? Can you use block nested loops join?
8. Describe how to evaluate a grouping query with aggregation operator MAX using a sorting-based approach.
9. Suppose that you are building a DBMS and want to add a new aggregate operator called SECOND LARGEST, which is a variation of the MAX operator. Describe how you would implement it.
10. Give an example of how buffer replacement policies can affect the performance of a join algorithm.

Exercise 14.2 Consider a relation R(*a, b, c, d, e*) containing 5,000,000 records, where each data page of the relation holds 10 records. R is organized as a sorted file with secondary indexes. Assume that *R.a* is a candidate key for R, with values lying in the range 0 to 4,999,999, and that R is stored in *R.a* order. For each of the following relational algebra queries, state which of the following approaches (or combination thereof) is most likely to be the cheapest:

- Access the sorted file for R directly.

- Use a clustered B+ tree index on attribute *R.a.*
- Use a linear hashed index on attribute *R.a.*
- Use a clustered B+ tree index on attributes *(R.a, R.b).*
- Use a linear hashed index on attributes *(R.a, R.b).*
- Use an unclustered B+ tree index on attribute *R.b.*

1. $O"u<50,OOOAb<50,ooo(R)$
2. $\sigma_{u=50,OOOAb<50,OOO}(R)$
3. $O"u>50,OOOAb=50,ooo(R)$
4. $\sigma_{u=50,OOOi\backslash a=50,OIO}(R)$
5. $\sigma_{a\neq 50,000\wedge b=50,000}(R)$
6. $\sigma_{a<50,OOOvb=50,OOO}(R)$

Exercise 14.3 Consider processing the following SQL projection query:

```
SELECT DISTINCT E.title, E.ename FROM Executives E
```

You are given the following information:

Executives has attributes *ename, title, dname,* and *address;* all are string fields of the same length.
The *ename* attribute is a candidate key.
The relation contains 10,000 pages,
There are 10 buffer pages.

Consider the optimized version of the sorting-based projection algorithm: The initial sorting pass reads the input relation and creates sorted runs of tuples containing only attributes *ename* and *title.* Subsequent merging passes eliminate duplicates while merging the initial runs to obtain a single sorted result (as opposed to doing a separate pass to eliminate duplicates from a sorted result containing duplicates).

1. How many sorted runs are produced in the first pass? What is the average length of these runs? (Assume that memory is utilized well and any available optimization to increase run size is used.) What is the I/O cost of this sorting pass?
2. How many additional merge passes are required to compute the final result of the projection query? What is the I/O cost of these additional passes?
3. (a) Suppose that a clustered B+ tree index on *title* is available. Is this index likely to offer a cheaper alternative to sorting? Would your answer change if the index were unclustered? Would your answer change if the index were a hash index?
   (b) Suppose that a clustered B+ tree index on *ename* is available. Is this index likely to offer a cheaper alternative to sorting? Would your answer change if the index were unclustered? Would your answer change if the index were a hash index?
   (c) Suppose that a clustered B+ tree index on *(ename, title)* is available. Is this index likely to offer a cheaper alternative to sorting? Would your answer change if the index were unclustered? Would your answer change if the index were a hash index?
4. Suppose that the query is as follows:

```
SELECT E.title, E.ename FROM Executives E
```

That is, you are not required to do duplicate elimination. How would your answers to the previous questions change?

Exercise 14.4 Consider the join $R \bowtie_{R.a=S.b} S$, given the following information about the relations to be joined. The cost metric is the number of page I/Os unless otherwise noted, and the cost of writing out the result should be uniformly ignored.

Relation R contains 10,000 tuples and has 10 tuples per page.
Relation S contains 2000 tuples and also has 10 tuples per page.
Attribute *b* of relation S is the primary key for S.
Both relations are stored as simple heap files.
Neither relation has any indexes built on it.
52 buffer pages are available.

1. What is the cost of joining R and S using a page-oriented simple nested loops join? What is the minimum number of buffer pages required for this cost to remain unchanged?
2. What is the cost of joining R and S using a block nested loops join? What is the minimum number of buffer pages required for this cost to remain unchanged?
3. What is the cost of joining R and S using a sort-merge join? What is the minimum number of buffer pages required for this cost to remain unchanged?
4. What is the cost of joining R and S using a hash join? What is the minimum number of buffer pages required for this cost to remain unchanged?
5. What would be the lowest possible I/O cost for joining R and S using *any* join algorithm, and how much buffer space would be needed to achieve this cost? Explain briefly.
6. How many tuples does the join of R and S produce, at most, and how many pages are required to store the result of the join back on disk?
7. Would your answers to any of the previous questions in this exercise change if you were told that *R.a* is a foreign key that refers to *S.b?*

Exercise 14.5 Consider the join of R and S described in Exercise 14.1.

1. With 52 buffer pages, if unclustered B+ indexes existed on *R.a* and *S.b*, would either provide a cheaper alternative for performing the join (using an index nested loops join) than a block nested loops join? Explain.
   (a) Would your answer change if only five buffer pages were available?
   (b) Would your answer change if S contained only 10 tuples instead of 2000 tuples?
2. With 52 buffer pages, if *clustered* B+ indexes existed on *R.a* and *S.b*, would either provide a cheaper alternative for performing the join (using the *index nested loops* algorithm) than a block nested loops join? Explain.
   (a) Would your answer change if only five buffer pages were available?
   (b) Would your answer change if S contained only 10 tuples instead of 2000 tuples?
3. If only 15 buffers were available, what would be the cost of a sort-merge join? What would be the cost of a hash join?
4. If the size of S were increased to also be 10,000 tuples, but only 15 buffer pages were available, what would be the cost of a sort-merge join? What would be the cost of a hash join?

5. If the size of S were increased to also be 10,000 tuples, and 52 buffer pages were available, what would be the cost of sort-merge join? What would be the cost of hash join?

Exercise 14.6 Answer each of the questions—if some question is inapplicable, explain why--in Exercise 14.1 again but using the following information about Rand S:

Relation R contains 200,000 tuples and has 20 tuples per page.
Relation S contains 4,000,000 tuples and also has 20 tuples per page.
Attribute *a* of relation R is the primary key for R.
Each tuple of R joins with exactly 20 tuples of S.
1,002 buffer pages are available.

Exercise 14.7 We described variations of the join operation called *outer joins* in Section 5.6.4 . One approach to implementing an outer join operation is to first evaluate the corresponding (inner) join and then add additional tuples padded with *null* values to the result in accordance with the semantics of the given outer join operator.. However, this requires us to compare the result of the inner join with the input relations to determine the additional tuples to be added. The cost of this comparison can be avoided by modifying the join algorithm to add these extra tuples to the result while input tuples are processed during the join. Consider the following join algorithms: *block nested loops join, index nested loops join, sort-merge join,* and *hash join.* Describe how you would modify each of these algorithms to compute the following operations on the Sailors and Reserves tables discussed in this chapter:

1. Sailors NATURAL LEFT OUTER JOIN Reserves
2. Sailors NATURAL RIGHT OUTER JOIN Reserves
3. Sailors NATURAL FULL OUTER JOIN Reserves

PROJECT-BASED EXERCISES

Exercise 14.8 *(Note to instructors: Additional details must be provided if this exen:ise* is *assigned; see Appendix 30.)* Implement the various join algorithms described in this chapter in Minibase. (As additional exercises, you Inay want to implement selected algorithms for the other operators as well.)

BIBLIOGRAPHIC NOTES

The implementation techniques used for relational operators in System R are discussed in [101]. The implementation techniques used in PRTV, which utilized relational algebra transformations and a form of multiple-query optimization, are discussed in [358]. The techniques used for aggregate operations in Ingres are described in [246]. [324] is an excellent survey of algorithms for implementing relational operators and is recommended for further reading.

Hash-based techniques are investigated (and compared with sort-based techniques) in [110], [222], [325], and [677]. Duplicate elimination is discussed in [99]. [277] discusses secondary storage access patterns arising in join implementations. Parallel algorithms for implementing relational operations are discussed in [99, 168, 220, 224, 233, 293,534].

15

A TYPICAL RELATIONAL QUERY OPTIMIZER

- How are SQL queries translated into relational algebra? As a consequence, what class of relation algebra queries does a query optimizer concentrate on?
- What information is stored in the system catalog of a DBMS and how is it used in query optimization?
- How does an optimizer estimate the cost of a query evaluation plan?
- How does an optimizer generate alternative plans for a query? What is the space of plans considered? What is the role of relational algebra equivalences in generating plans?
- How are nested SQL queries optimized?
- Key concepts: SQL to algebra, query block; system catalog, data dictionary, metadata, system statistics, relational representation of catalogs; cost estimation, size estimation, reduction factors; histograms, equiwidth, equidepth, compressed; algebra equivalences, pushing selections, join ordering; plan space, single-relation plans, multi-relation left-deep plans; enumerating plans, dynamic programming approach, alternative approaches

Life is what happens while you're busy making other plam.

-John Lennon

In this chapter, we present a typical relational query optimizer in detail. We begin by discussing how SQL queries are converted into units called *blocks*

and how blocks are translated into (extended) relational algebra expressions (Section 15.1). The central task of an optimizer is to find a good plan for evaluating such expressions. Optimizing a relational algebra expression involves two basic steps:

- Enumerating alternative plans for evaluating the expression. Typically, an optimizer considers a subset of all possible plans because the number of possible plans is very large.
- Estimating the cost of each enumerated plan and choosing the plan with the lowest estimated cost.

We discuss how to use system statistics to estimate the properties of the result of a relational operation, in particular result sizes, in Section 15.2. After discussing how to estimate the cost of a given plan, we describe the space of plans considered by a typical relational query optimizer in Sections 15.3 and 15.4. We discuss how nested SQL queries are handled in Section 15.5. We briefly discuss some of the influential choices made in the System R query optimizer in Section 15.6. We conclude with a short discussion of other approaches to query optimization in Section 15.7.

We consider a number of example queries using the following schema:

Sailors(sid: integer, *sname:* string, *rating:* integer, *age:* real)
Boats(*bid:* integer, *bname:* string, *color:* string)
Reserves(sid: integer, *bid:* integer, *day:* dates, *mame:* string)

As in Chapter 14, we assume that each tuple of Reserves is 40 bytes long, that a page can hold 100 Reserves tuples, and that we have 1000 pages of such tuples. Similarly, we assume that each tuple of Sailors is 50 bytes long, that a page can hold 80 Sailors tuples, and that we have 500 pages of such tuples.

15.1 TRANSLATING SQL QUERIES INTO ALGEBRA

SQL queries are optimized by decomposing them into a collection of smaller units, called *blocks*. A typical relational query optimizer concentrates on optimizing a single block at a time. In this section, we describe how a query is decomposed into blocks and how the optimization of a single block can be understood in tenus of plans composed of relational algebra operators.

15.1.1 Decomposition of a Query into Blocks

When a user submits an SQL query, the query is parsed into a collection of query blocks and then passed on to the query optimizer. A **query block**

```
SELECT    S.siel, MIN (Relay)
FROM      Sailors S, Reserves R, Boats B
WHERE     S.siel = Rsiel AND Rbid = B.bid AND Rcolor = 'red' AND
          S.rating = ( SELECT MAX (S2.rating)
                       FROM   Sailors S2 )
GROUP BY  S.sid
HAVING    COUNT (*) > 1
```

Figure 15.1 Sailors Reserving Red Boats

(or simply block) is an SQL query with no nesting and exactly one SELECT clause and one FROM clause and at most one WHERE clause, GROUP BY clause, and HAVING clause. The WHERE clause is assumed to be in conjunctive normal form, as per the discussion in Section 14.2. We use the following query as a running example:

For each 8ailor' with the highe8t mting (oveT all sailors) and at least two reservations for Ted boats, find the sailoT id and the earliest date on which the sailor has a TeseTvat:ion for a red boat.

The SQL version of this query is shown in Figure 15.1. This query has two query blocks. The nested block is:

```
SELECT MAX (S2.rating)
FROM   Sailors S2
```

The nested block computes the highest sailor rating. The outer block is shown in Figure 15.2. Every SQL query can be decomposed into a collection of query blocks without nesting.

```
SELECT    S.sid, MIN (Rday)
FROM      Sailors S, Reserves R, Boats B
WHERE     S.sid = Rsiel AND Rbicl = B.bid AND Rcolor = 'red' AND
          S.rating = RefeTence to nested block
GROUP BY  S.sid
HAVING    COUNT (*) > 1
```

Figure 15.2 Outer Block of Red Boats Query

The optimizer examines the system catalogs to retrieve information about the types and lengths of fields, statistics about the referenced relations, and the access paths (indexes) available for them. The optimizer then considers each query block and chooses a query evaluation plan for that block. We focus 1nostly on optimizing a single query block and defer a discussion of nested queries to Section 15.5.

15.1.2 A Query Block as a Relational Algebra Expression

The first step in optimizing a query block is to express it as a relational algebra expression. For uniformity, let us assume that GROUP BY and HAVING are also operators in the extended algebra used for plans and that aggregate operations are allowed to appear in the argument list of the projection operator. The meaning of the operators should be clear from our discussion of SQL. The SQL query of Figure 15.2 can be expressed in the extended algebra as:

$$\pi_{S.sid,MIN(R.day)}($$
$$HAVINGcoUNT(*»2($$
$$GROUP\ BYs._{sid}($$
$$\sigma_{S.sid=R.sid\wedge R.bid=B.bidAB.coloT='Ted'AS.rating=value\_from\_nested\_block}($$
$$Sailors \times Reserves \times Boats))))$$

For brevity, we used *S*, *R*, and *B* (rather than Sailors, Reserves, and Boats) to prefix attributes. Intuitively, the selection is applied to the cross-product of the three relations. Then the qualifying tuples are grouped by *S.sid*, and the HAVING clause condition is used to discard some groups. For each remaining group, a result tuple containing the attributes (and count) mentioned in the projection list is generated. This algebra expression is a faithful summary of the semantics of an SQL query, which we discussed in Chapter 5.

Every SQL query block can be expressed as an extended algebra expression having this form. The SELECT clause corresponds to the projection operator, the WHERE clause corresponds to the selection operator, the FROM clause corresponds to the cross-product of relations, and the remaining clauses are mapped to corresponding operators in a straightforward manner.

The alternative plans examined by a typical relational query optimizer can be understood by recognizing that *a query is essentially treated as a* $\sigma\pi\times$ *algebm expression*, with the remaining operations (if any, in a given query) carried out on the result of the $\sigma\pi\times$ expression. The $\sigma\pi\times$ expression for the query in Figure 15.2 is:

$$\pi_{S.sid,R.day}($$
$$\sigma_{S.sid=R.sidAR.b'id=B.bid\wedge B.color='red'\wedge S.rating=value\_fTO/lLTlcste(Lblock}($$
$$Sailors \times Reserves \times Boats))$$

To make sure that the GROUP BY and HAVING operations in the query can be carried out, the attributes mentioned in these clauses are added to the projection list. Further, since aggregate operations in the SELECT cla.use, such as the MIN *(R.day)* operation in our example, are computed after first computing the $\sigma\pi\times$ part of the query, aggregate expressions in the projectioll list are replaced

by the names of the attributes to which they refer. Thus, the optimization of the $\sigma\pi$x part of the query essentially ignores these aggregate operations.

The optimizer finds the best plan for the $\sigma\pi$x expression obtained in this manner from a query. This plan is evaluated and the resulting tuples are then sorted (alternatively, hashed) to implement the GROUP BY clause. The HAVING clause is applied to eliminate some groups, and aggregate expressions in the SELECT clause are computed for each remaining group. This procedure is summarized in the following extended algebra expression:

$\pi_{S.sid,MIN(R.day)}$ (
HAVING$_{COUNT(*)>2}$(
GROUP BY$_{S.sid}$(
$\pi_{S.sid,R.day}$(
$\sigma_{S.sid=R.sid \wedge R.bid=B.bid \wedge B.color='red' \wedge S.rating=value\_from\_nested\_block}$(
Sailors x *Reserves* x *Boats*)))))

Some optimizations are possible if the FROM clause contains just one relation and the relation has some indexes that can be used to carry out the grouping operation. We discuss this situation further in Section 15.4.1.

To a first approximation therefore, the alternative plans examined by a typical optimizer can be understood in terms of the plans considered for $\sigma\pi$x queries. An optimizer enumerates plans by applying several equivalences between relational algebra expressions, which we present in Section 15.3. We discuss the space of plans enumerated by an optimizer in Section 15.4.

15.2 ESTIMATING THE COST OF A PLAN

For each enumerated plan, we have to estimate its cost. There are two parts to estimating the cost of an evaluation plan for a query block:

1. For each node in the tree, we must *estimate the cost* of performing the corresponding operation. Costs are affected significantly by whether pipelining is used or temporary relations are created to pass the output of an operator to its parent.
2. For each node in the tree, we must *estimate the size of the result* and whether it is sorted. This result is the input for the operation that corresponds to the parent of the current node, and the size and sort order in turn affect the estimation of size, cost, and sort order for the parent.

We discussed the cost of implementation techniques for relational operators in Chapter 14. As we saw there, estimating costs requires knowledge of various

parameters of the input relations, such as the number of pages and available indexes. Such statistics are maintained in the DBMS's system catalogs. In this section, we describe the statistics maintained by a typical DBMS and discuss how result sizes are estimated. As in Chapter 14, we use the number of page I/Os as the metric of cost and ignore issues such as blocked access, for the sake of simplicity.

The estimates used by a DBMS for result sizes and costs are at best approximations to actual sizes and costs. It is unrealistic to expect an optimizer to find the very best plan; it is more important to avoid the worst plans and find a good plan.

15.2.1 Estimating Result Sizes

We now discuss how a typical optimizer estimates the size of the result computed by an operator on given inputs. Size estimation plays an important role in cost estimation as well because the output of one operator can be the input to another operator, and the cost of an operator depends on the size of its inputs.

Consider a query block of the form:

SELECT *attTibute list*
FROM *Telation list*
WHERE *teTml 1\ teTm2 1\ . . . 1\ teTm$_n$*

The maximum number of tuples in the result of this query (without duplicate elimination) is the product of the cardinalities of the relations in the FROM clause. Every term in the WHERE clause, however, eliminates some of these potential result tuples. We can model the effect of the WHERE clause on the result size by associating a **reduction factor** with each term, which is the ratio of the (expected) result size to the input size considering only the selection represented by the term. The actual size of the result can be estimated as the maximum size times the product of the reduction factors for the terms in the WHERE clause. Of course, this estimate reflects the unrealistic but simplifying assumption that the conditions tested by each term are statistically independent.

We now consider how reduction factors can be computed for different kinds of terms in the WHERE clause by using the statistics available in the catalogs:

- *column* = *value*: For a term of this form, the reduction factor can be approximated by $\frac{1}{Keys(I)}$ if there is an index I on *column* for the relation in question. This formula assumes uniform distribution of tuples among the

index key values; this uniform distribution assumption is frequently made in arriving at cost estimates in a typical relational query optimizer. If there is no index on *col'umn,* the System R optimizer arbitrarily assumes that the reduction factor is $\frac{1}{10}$. Of course, it is possible to maintain statistics such as the number of distinct values present for any attribute whether or not there is an index on that attribute. If such statistics are maintained, we can do better than the arbitrary choice of $\frac{1}{10}$.

- *column1 = column2:* In this case the reduction factor can be approximated by $\frac{1}{MAX(NKeys(I1),NKeys(I2))}$ if there are indexes I1 and 12 on *column1* and *column2,* respectively. This formula assumes that each key value in the smaller index, say, I1, has a matching value in the other index. Given a value for *columnl,* we assume that each of the $NKeys(I2)$ values for *column2* is equally likely. Therefore, the number of tuples that have the same value in *column2* as a given value in *columni* is $\frac{1}{NKeys(I2)}$. If only one of the two columns has an index *I,* we take the reduction factor to be $\frac{1}{NKeys(I)}$; if neither column has an index, we approximate it by the ubiquitous $\frac{1}{10}$. These formulas are used whether or not the two columns appear in the same relation.
- *column > value:* The reduction factor is approximated by $\frac{High(I) - value}{High(I) - Low(I)}$ if there is an index I on *column.* If the column is not of an arithmetic type or there is no index, a fraction less than half is arbitrarily chosen. Similar formulas for the reduction factor can be derived for other range selections.
- *column* IN *(list of values):* The reduction factor is taken to be the reduction factor for *column = value* multiplied by the number of items in the list. However, it is allowed to be at most half, reflecting the heuristic belief that each selection eliminates at least half the candidate tuples.

These estimates for reduction factors are at best approximations that rely on assumptions such as uniform distribution of values and independent distribution of values in different columns. In recent years more sophisticated techniques based on storing more detailed statistics (e.g., histograms of the values in a column, which we consider later in this section) have been proposed and are finding their way into commercial systems.

Reduction factors can also be approximated for terms of the form *column* IN *subquery* (ratio of the estimated size of the subquery result to the number of distinct values in *column* in the outer relation); NOT *condition* (I-reduction factor for *condition*); *value1<column<value2;* the disjunction of two conditions; and so on, but we will not discuss such reduction factors.

To summarize, regardless of the plan chosen, we can estimate the size of the final result by taking the product of the sizes of the relations in the FROM clause

Estimating Query Characteristics: IBM DB2, Informix, Microsoft SQL Server, Oracle 8, and Sybase ASE all usehistograms to estimate query characteristics such as result size and cost. As an example, Sybase ASE uses one-dimensional, equidepth histograms with some special attention paid to high frequency values, so that their count is estimated accurately. ASE also keeps the average count of duplicates for each prefix of an index to estimate correlations between histograms for composite keys (although it does not maintain such histograms). ASE also maintains estimates of the degree of clustering in tables and indexes. IBM DB2, Informix, and Oracle also use one-dimensional equidepth histograms; Oracle automatically switches to maintaining a count of duplicates for each value when there are few values in a column. Microsoft SQL Server uses one-dimensional equiarea histograms with some optimizations (adjacent buckets with similar distributions are sometimes combined to compress the histogram). In SQL Server, the creation and maintenance of histograms is done automatically with no need for user input.

Although sampling techniques have been studied for estimating result sizes and costs, in current systems, sampling is used only by system utilities to estimate statistics or build histograms but not directly by the optimizer to estimate query characteristics. Sometimes, sampling is used to do load balancing in parallel implementations.

and the reduction factors for the terms in the WHERE clause. We can similarly estimate the size of the result of each operator in a plan tree by using reduction factors, since the subtree rooted at that operator's node is itself a query block.

Note that the number of tuples in the result is not affected by projections if duplicate elimination is not performed. However, projections reduce the number of pages in the result because tuples in the result of a projection are smaller than the original tuples; the ratio of tuple sizes can be used as a reduction factor for projection to estimate the result size in pages, given the size of the input relation.

Improved Statistics: Histograms

Consider a relation with N tuples and a selection of the form *colu:rnn* > *value* on a column with an index *I*. The reduction factor r is approximated by $\frac{High(I) - value}{High(I) - Low(I)}$ and the size of the result is estimated as *TN*. This estimate relies on the assumption that the distribution of values is uniform.

Estimates can be improved considerably by maintaining more detailed statistics than just the low and high values in the index *I*. Intuitively, we want to approximate the distribution of key values *I* as accurately as possible. Consider the two distributions of values shown in Figure 15.3. The first is a nonuniform distribution *D* of values (say, for an attribute called *age).* The *frequency* of a value is the number of tuples with that *age* value; a distribution is represented by showing the frequency for each possible *age* value. In our example, the lowest *age* value is 0, the highest is 14, and all recorded *age* values are integers in the range 0 to 14. The second distribution approximates *D* by assuming that each *age* value in the range a to 14 appears equally often in the underlying collection of tuples. This approximation can be stored compactly because we need to record only the low and high values for the *age* range (0 and 14 respectively) and the total count of all frequencies (which is 45 in our example).

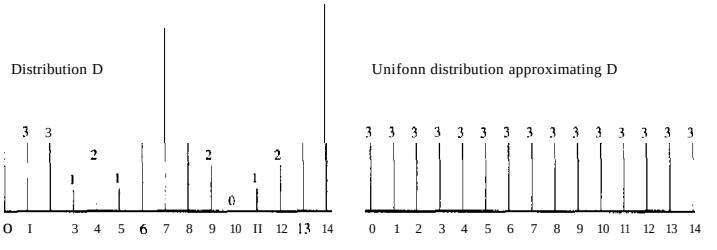


Figure 15.3 Uniform vs. Nonuniform Distributions

Consider the selection $age > 13$. From the distribution *D* in Figure 15.3, we see that the result has 9 tuples. Using the uniform distribution approximation, on the other hand, we estimate the result size as $\frac{1}{15} \cdot 45 = 3$ tuples. Clearly, the estimate is quite inaccurate.

A histogram is a data structure maintained by a DBMS to approximate a data distribution. In Figure 15.4, we show how the data distribution from Figure 15.3 can be approximated by dividing the range of *age* values into subranges called buckets, and for each bucket, counting the number of tuples with *age* values within that bucket. Figure 15.4 shows two different kinds of histograms, called *equiwidth* and *equidepth,* respectively.

Consider the selection query $age > 13$ again and the first (equiwidth) histogram. We can estimate the size of the result to be 5 because the selected range includes a third of the range for Bucket 5. Since Bucket 5 represents a total of 15 tuples, the selected range corresponds to $\frac{1}{3} \cdot 15 = 5$ tuples. As this example shows, we a..ssume that the distribution *within* a histogram bucket is uniform. Therefore, when we simply maintain the high and low values for index

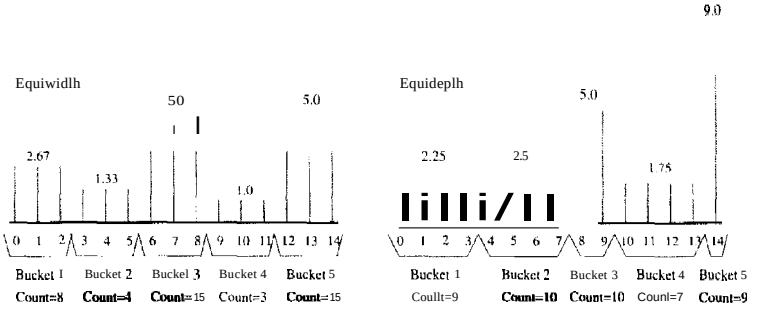


Figure 15.4 Histograms Approximating Distribution *D*

I, we effectively use a 'histogram' with a single bucket. Using histograms with a small number of buckets instead leads to much more accurate estimates, at the cost of a few hundred bytes per histogram. (Like all statistics in a DBMS, histograms are updated periodically rather than whenever the data is changed.)

One important question is how to divide the value range into buckets. In an equiwidth histogram, we divide the range into subranges of equal size (in terms of the *age* value range). We could also choose subranges such that the number of tuples within each subrange (i.e., bucket) is equal. Such a histogram, called an equidepth histogram, is also illustrated in Figure 15.4. Consider the selection *age* > 13 again. Using the equidepth histogram, we are led to Bucket 5, which contains only the *age* value 15, and thus we arrive at the exact answer, 9. While the relevant bucket (or buckets) generally contains more than one tuple, equidepth histograms provide better estimates than equiwidth histograms. Intuitively, buckets with very frequently occurring values contain fewer values, and thus the uniform distribution assumption is applied to a smaller range of values, leading to better approximations. Conversely, buckets with mostly infrequent values are approximated less accurately in an equidepth histogram, but for good estimation, the frequent values are important.

Proceeding further with the intuition about the importance of frequent values, another alternative is to maintain separate counts for a small number of very frequent values, say the *age* values 7 and 14 in our example, and maintain an equidepth (or other) histogram to cover the remaining values. Such a histogram is called a compressed histogram. Most commercial DB1\1Ss currently use equidepth histograms, and some use compressed histograms.

15.3 RELATIONAL ALGEBRA EQUIVALENCES

In this section, we present several equivalences among relational algebra expressions; and in Section 15.4, we discuss the space of alternative plans considered by a optimizer.

Our discussion of equivalences is aimed at explaining the role that such equivalences play in a System R style optimizer. In essence, a basic SQL query block can be thought of as an algebra expression consisting of the cross-product of all relations in the FROM clause, the selections in the WHERE clause, and the projections in the SELECT clause. The optimizer can choose to evaluate any equivalent expression and still obtain the same result. Algebra equivalences allow us to convert cross-products to joins, choose different join orders, and push selections and projections ahead of joins. For simplicity, we assume that naming conflicts never arise and we need not consider the renaming operator ρ.

15.3.1 Selections

Two important equivalences involve the selection operation. The first one involves cascading of selections:

$$\sigma_{c_1 \wedge c_2 \wedge \ldots c_n}(R) \equiv \sigma_{c_1}(\sigma_{c_2}(\ldots(\sigma_{c_n}(R))\ldots))$$

Going from the right side to the left, this equivalence allows us to combine several selections into one selection. Intuitively, we can test whether a tuple meets each of the conditions $c_1 \ldots c_n$ at the same time. In the other direction, this equivalence allows us to take a selection condition involving several conjuncts and replace it with several smaller selection operations. Replacing a selection with several smaller selections turns out to be very useful in combination with other equivalences, especially commutation of selections with joins or cross-products, which we discuss shortly. Intuitively, such a replacement is useful in cases where only part of a complex selection condition can be pushed.

The second equivalence states that selections are commutative:

$$\sigma_{c_1}(\sigma_{c_2}(R)) \equiv \sigma_{c_2}(\sigma_{c_1}(R))$$

In other words,' we can test the conditions c_1 and c_2 in either order.

15.3.2 Projections

The rule for cascading projections says that successively elilninating columns from a relation is equivalent to sirnply eliminating all but the columns retained

by the final projection:

$$\pi_{a_1}(R) \equiv \pi_{a_1}(\pi_{a_2}(\ldots(\pi_{a_n}(R))\ldots))$$

Each a_i is a set of attributes of relation R, and $ai \subseteq a_{i+1}$ for i = 1 ... n - 1. This equivalence is useful in conjunction with other equivalences such as commutation of projections with joins.

15.3.3 Cross-Products and Joins

Two important equivalences involving cross-products and joins. We present them in terms of natural joins for simplicity, but they hold for general joins as well.

First, assuming that fields are identified by name rather than position, these operations are commutative:

$$Rx8 \qquad 8xR$$

$$RN8$$

This property is very important. It allows us to choose which relation is to be the inner and which the outer in a join of two relations.

The second equivalence states that joins and cross-products are associative:

$$R \times (8 \times T) \qquad (R \times 8) \times T$$

$$RN\,(8NT) \qquad (R \bowtie 8)\,NT$$

Thus we can either join Rand 8 first and then join T to the result, or join 8 and T first and then join R to the result. The intuition behind associativity of cross-products is that, regardless of the order in which the three relations are considered, the final result contains the same columns. Join associativity is based on the same intuition, with the additional observation that the selections specifying the join conditions can be cascaded. Thus the same rows appear in the final result, regardless of the order in which the relations are joined.

Together with commutativity, associativity essentially says that we can choose to join any pair of these relations, then join the result with the third relation, and always obtain the same final result. For example, let us verify that

$$R \bowtie (8 \bowtie T) \equiv (T \bowtie R) \bowtie 8$$

From commutativity, we have:

$$RN\,(8NT) \qquad RN\,(TN\,8)$$

From associativity, we have:

$$R \bowtie (T \bowtie S) \qquad (R \bowtie T) \bowtie S$$

Using commutativity again, we have:

$$(R \bowtie T) \bowtie S \equiv (T \bowtie R) \bowtie S$$

In other words, when joining several relations, we are free to join the relations in any order we choose. This order-independence is fundamental to how a query optimizer generates alternative query evaluation plans.

15.3.4 Selects, Projects, and Joins

Some important equivalences involve two or more operators.

We can **commute** a selection with a projection if the selection operation involves only attributes retained by the projection:

$$\pi_a(\sigma_c(R)) \equiv \sigma_c(\pi_a(R))$$

Every attribute mentioned in the selection condition c must be included in the set of attributes *a*.

We can **combine** a selection with a cross-product to form a join, as per the definition of join:

$$R \bowtie_c S \equiv \sigma_c(R \times S)$$

We can **commute** a selection with a cross-product or a join if the selection condition involves only attributes of one of the arguments to the cross-product or join:

$$\sigma_c(R \text{ x } S) \qquad \sigma_c(R) \text{ x } S$$
$$\sigma_c(R \bowtie S) \qquad \sigma_c(R) \bowtie S$$

The attributes mentioned in c must appear only in *R* and not in *S*. Similar equivalences hold if c involves only attributes of *S* and not *R*, of course.

In general, a selection σ_c on *R* x *S* can be replaced by a cascade of selections σ_{c_1}, σ_{c_2}, and σ_{c_3} such that c1 involves attributes of both R and *S*, c2 involves only attributes of *R*, and c_3 involves only attributes of *S*:

$$\sigma_c(R \times S) \equiv \sigma_{c_1 \wedge c_2 \wedge c_3}(R \times S)$$

Using the cascading rule for selections, this expression is equivalent to

$$\sigma_{c_1}(\sigma_{c_2}(\sigma_{c_3}(R \times S)))$$

Using the rule for commuting selections and cross-products, this expression is equivalent to

$$\sigma_{C1}(\sigma_{C2}(R) \times \sigma_{C3}(S))$$

Thus we can push part of the selection condition c ahead of the cross-product. This observation also holds for selections in combination with joins. of course.

We can **commute** a projection with a cross-product:

$$\pi_a(R \times S) \equiv \pi_{a_1}(R) \times \pi_{a_2}(S)$$

where a1 is the subset of attributes in *a* that appear in *R*, and *a2* is the subset of attributes in *a* that appear in *S*. We can also **commute** a projection with a join if the join condition involves only attributes retained by the projection:

$$\pi_a(R \bowtie_c S) \equiv \pi_{a_1}(R) \bowtie_c \pi_{a_2}(S)$$

where a1 is the subset of attributes in *a* that appear in *R*, and *a2* is the subset of attributes in *a* that appear in *S*. Further, every attribute mentioned in the join condition c must appear in *a*.

Intuitively, we need to retain only those attributes of R and *S* that are either mentioned in the join condition c or included in the set of attributes *a* retained by the projection. Clearly, if *a* includes all attributes mentioned in c, the previous commutation rules hold. If *a* does *not* include all attributes mentioned in c, we can generalize the commutation rules by first projecting out attributes that are not mentioned in c or *a*, performing the join, and then projecting out all attributes that are not in *a*:

$$\pi_a(R \bowtie_c S) \equiv \pi_a(\pi_{a_1}(R) \bowtie_c \pi_{a_2}(S))$$

Now, a_1 is the subset of attributes of *R* that appear in either *a* or c, and a_2 is the subset of attributes of *S* that appear in either *a* or c.

We can in fact derive the more general commutation rule by using the rule for cascading projections and the simple commutation rule, and we leave this as an exercise for the reader.

15.3.5 Other Equivalences

Additional equivalences hold when we consider operations such as set-difference, union, and intersection. Union and intersection are associative and commutative. Selections and projections can be commuted with each of the set operations (set-difference, union, and intersection). We do not discuss these equivalences further.

```
SELECT    S.rating, COUNT (*)
FROM      Sailors S
WHERE     S.rating > 5 AND S.age = 20
GROUP BY  S.rating
HAVING    COUNT DISTINCT (S.sname) > 2
```

Figure 15.5 A Single-Relation Query

15.4 ENUMERATION OF ALTERNATIVE PLANS

We now come to an issue that is at the heart of an optimizer, namely, the space of alternative plans considered for a given query. Given a query, an optimizer essentially enumerates a certain set of plans and chooses the plan with the least estimated cost; the discussion in Section 12.1.1 indicated how the cost of a plan is estimated. The algebraic equivalences discussed in Section 15.3 form the basis for generating alternative plans, in conjunction with the choice of implementation technique for the relational operators (e.g., joins) present in the query. However, not all algebraically equivalent plans are considered, because doing so would make the cost of optimization prohibitively expensive for all but the simplest queries. This section describes the subset of plans considered by a typical optimizer.

There are two important cases to consider: queries in which the FROM clause contains a single relation and queries in which the FROM clause contains two or more relations.

15.4.1 Single-Relation Queries

If the query contains a single relation in the FROM clause, only selection, projection, grouping, and aggregate operations are involved; there are no joins. If we have just one selection or projection or aggregate operation applied to a relation, the alternative implementation techniques and cost estimates discussed in Chapter 14 cover all the plans that must be considered. We now consider how to optimize queries that involve a combination of several such operations, using the following query as an example:

For each rating greater than 5, print the rating and the number of 20-year'-old sailors with that rating, provided that there are at least two such sailors with different names.

The SQL version of this query is shown in Figure 15.5. Using the extended algebra notation introduced in Section 15.1.2, we can write this query as:

$$\pi_{S.rating, COUNT(*)}($$

$$HAVING_{COUNT\,DISTINCT(S.sname)>2}($$
$$GROUP\ BY_{S.rating}($$
$$\pi_{S.rating,S.sname}($$
$$\sigma_{S.rating>5 \wedge S.age=20}($$
$$Sailors)))))$$

Notice that *S.sname* is added to the projection list, even though it is not in the SELECT clause, because it is required to test the HAVING clause condition.

We are now ready to discuss the plans that an optimizer would consider. The main decision to be made is which access path to use in retrieving Sailors tuples. If we considered only the selections, we would simply choose the most selective access path, based on which available indexes *match* the conditions in the WHERE clause (as per the definition in Section 14.2.1). Given the additional operators in this query, we must also take into account the cost of subsequent sorting steps and consider whether these operations can be performed without sorting by exploiting some index. We first discuss the plans generated when there are no suitable indexes and then examine plans that utilize some index.

Plans without Indexes

The basic approach in the absence of a suitable index is to scan the Sailors relation and apply the selection and projection (without duplicate elimination) operations to each retrieved tuple, as indicated by the following algebra expression:

$$\pi_{S.rating,S.sname}($$
$$\sigma_{S.rating>5 \wedge S.age=20}($$
$$Sailors))$$

The resulting tuples are then sorted according to the GROUP BY clause (in the example query, on *rating*), and one answer tuple is generated for each group that meets the condition in the HAVING clause. The computation of the aggregate functions in the SELECT and HAVING clauses is done for each group, using one of the techniques described in Section 14.6.

The cost of this approach consists of the costs of each of these steps:

1. Perfonning a file scan to retrieve tuples and apply the selections and projections.
2. 'Writing out tuples after the selections and projectiolls.
3. Sorting these tuples to implement the GROUP BY clause.

Note that the HAVING clause does not cause additional I/O. The aggregate computations can be done on-the-fiy (with respect to I/O) as we generate the tuples in each group at the end of the sorting step for the GROUP BY clause.

In the example query the cost includes the cost of a file scan on Sailors plus the cost of writing out *(S.rating, S.sname)* pairs plus the cost of sorting as per the GROUP BY clause. The cost of the file scan is *NPages(Sailors)*, which is 500 I/Os, and the cost of writing out *(S.rating, S.sname)* pairs is *NPages(Sailors)* times the ratio of the size of such a pair to the size of a Sailors tuple times the reduction factors of the two selection conditiolls. In our example, the result tuple size ratio is about 0.8, the *rating* selection has a reduction factor of 0.5, and we use the default factor of 0.1 for the *age* selection. Therefore, the cost of this step is 20 l/Os. The cost of sorting this intermediate relation (which we call *Temp)* can be estimated as *3*NPages(Temp)*, which is 60 I/Os, if we assume that enough pages are available in the buffer pool to sort it in two passes. (Relational optimizers often assume that a relation can be sorted in two passes, to simplify the estimation of sorting costs. If this assumption is not met at run-time, the actual cost of sorting may be higher than the estimate.) The total cost of the example query is therefore 500 + 20 + 60 = 580 l/Os.

Plans Utilizing an Index

Indexes can be utilized in several ways and can lead to plans that are significantly faster than any plan that does not utilize indexes:

1. Single-Index Access Path: If several indexes match the selection conditions in the WHERE clause, each matching index offers an alternative access path. An optimizer can choose the access path that it estimates will result in retrieving the fewest pages, apply any projections and nonprimary selection terms (i.e., parts of the selection condition that do not match the index), and proceed to compute the grouping and aggregation operations (by sorting on the GROUP BY attributes).

2. Multiple-Index Access Path: If several indexes using Alternatives (2) or (3) for data entries match the selection condition, each such index can be used to retrieve a set of rids. We can *intersect* these sets of rids, then sort the result by page id (assuming that the rid representation includes the page id) and retrieve tuples that satisfy the primary selection terms of all the matching indexes. Any projections and nonprimary selection terms can then be applied, followed by grouping and aggregation operations.

3. Sorted Index Access Path: If the list of grouping attributes is a prefix of a trec index, the index can be used to retrieve tuples in the order required by the GROUP BY clause. All selection conditions can be applied on each

retrieved tuple, unwanted fields can be removed, and aggregate operations computed for each group. This strategy works well for clustered indexes.

4. Index-Only Access Path: If all the attributes mentioned in the query (in the SELECT, WHERE, GROUP BY, or HAVING clauses) are included in the search key for some *dense* index on the relation in the FROM clause, an index-only scan can be used to compute answers. Because the data entries in the index contain all the attributes of a tuple needed for this query and there is one index entry per tuple, we never need to retrieve actual tuples from the relation. Using just the data entries from the index, we can carry out the following steps as needed in a given query: Apply selection conditions, remove unwanted attributes, sort the result to achieve grouping, and compute aggregate functions within each group. This *index-only* approach works even if the index does not match the selections in the WHERE clause. If the index matches the selection, we need examine only a subset of the index entries; otherwise, we must scan all index entries. In either case, we can avoid retrieving actual data records; therefore, the cost of this strategy does not depend on whether the index is clustered. In addition, if the index is a tree index and the list of attributes in the GROUP BY clause forms a prefix of the index key, we can retrieve data entries in the order needed for the GROUP BY clause and thereby avoid sorting!

We now illustrate each of these four cases, using the query shown in Figure 15.5 as a running example. We assume that the following indexes, all using Alternative (2) for data entries, are available: a B+ tree index on *rating*, a hash index on *age*, and a B+ tree index on *(rating, sname, age)*. For brevity, we do not present detailed cost calculations, but the reader should be able to calculate the cost of each plan. The steps in these plans are scans (a file scan, a scan retrieving tuples by using an index, or a scan of only index entries), sorting, and writing temporary relations; and we have already discussed how to estimate the costs of these operations.

As an example of the first case, we could choose to retrieve Sailors tuples such that *S.age=20* using the hash index on *age*. The cost of this step is the cost of retrieving the index entries plus the cost of retrieving the corresponding Sailors tuples, which depends on whether the index is clustered. We can then apply the condition *S.rating* > 5 to each retrieved tuple; project out fields not mentioned in the SELECT, GROUP BY, and HAVING clauses; and write the result to a temporary relation. In the example, only the *rating* and *sname* fields need to be retained. The temporary relation is then sorted on the *rating* field to identify the groups, and some groups are eliminated by applying the HAVING conclitioIl.

Utilizing Indexes: All of the main RDBMSs recognize the importance of index-only plans and look for such plans whenever possible. In IBM DD2, when creating an index a user can specify a set of 'include' columns that are to be kept in the index but are *not* part of the index key. This allows a richer set of index-only queries to be handled, because columns frequently accessed are included in the index even if they are ;notpart of the key. In Microsoft SQL Server, an interesting class of index-only plans is considered: Consider a query that selects attributes *sal* and *age* from a table, given an index on *sal* and another index on *age*. SQL Server uses the indexes by joining the entries on the rid of data records to identify *(sal, age)* pairs that appear in the table.

As an example of the second case, we can retrieve rids of tuples satisfying *mting>5* using the index on *rating*, retrieve rids of tuples satisfying *age=20* using the index on *age*, sort the retrieved rids by page number, and then retrieve the corresponding Sailors tuples. We can retain just the *rating* and *name* fields and write the result to a temporary relation, which we can sort on *rating* to implement the GROUP BY clause. (A good optimizer might pipeline the projected tuples to the sort operator without creating a temporary relation.) The HAVING clause is handled as before.

As an example of the third case, we can retrieve Sailors tuples in which *S.mting* > 5, ordered by *rating*, using the B+ tree index on *rating*. We can compute the aggregate functions in the HAVING and SELECT clauses on-the-fly because tuples are retrieved in *rating* order.

As an example of the fourth case, we can retrieve *data entries* from the *(mting, sname, age)* index in which *mting* > 5. These entries are sorted by *rating* (and then by *snarne* and *age*, although this additional ordering is not relevant for this query). We can choose entries with *age=20* and compute the aggregate functions in the HAVING and SELECT clauses on-the-fly because the data entries are retrieved in *rating* order. In this case, in contrast to the previous case, we do not retrieve any Sailors tuples. This property of not retrieving data records makes the index-only strategy especially valuable with unclusterecl indexes.

15.4.2 Multiple-Relation Queries

Query blocks that contain two or more relations in the FROM clause require joins (or cross-products). Finding a good plan for such queries is very important because these queries can be quite expensive. Regardless of the plan chosen, the size of the final result can be estimated by taking the product of the sizes

of the relations in the FROM clause and the reduction factors for the terms in the WHERE clause. But, depending on the order in which relations are joined, intermediate relations of widely varying sizes can be created, leading to plans with very different costs.

Enumeration of Left-Deep Plans

As we saw in Chapter 12, current relational systems, following the lead of the System R optimizer, only consider left-deep plans. We now discuss how this dass of plans is efficiently searched using dynamic programming.

Consider a query block of the form:

SELECT *attribute list*
FROM *relation list*
WHERE $term_1$ 1\ *term2* 1\ ... 1\ $term_n$

A System R style query optimizer enumerates all left-deep plans, with selections and projections considered (but not necessarily applied!) as early as possible. The enumeration of plans can be understood as a multiple-pass algorithm in which we proceed as follows:

Pass 1: We enumerate all single-relation plans (over some relation in the FROM clause). Intuitively, each single-relation plan is a partial left-deep plan for evaluating the query in which the given relation is the first (in the linear join order for the left-deep plan of which it is a part). When considering plans involving a relation *A*, we identify those selection terms in the WHERE clause that mention only attributes of *A*. These are the selections that can be performed when first accessing *A*, before any joins that involve *A*. We also identify those attributes of *A* not mentioned in the SELECT clause or in terms in the WHERE clause involving attributes of other relations. These attributes can be projected out when first accessing *A*, before any joins that involve *A*. We choose the best access method for *A* to carry out these selections and projections, as per the discussion in Section 15.4.1.

For each relation, if we find plans that produce tuples in different orders, we retain the cheapest plan for each such ordering of tuples. An ordering of tuples could prove useful at a subsequent step, say, for a sort-merge join or implementing a GROUP BY or ORDER BY clause. Hence, for a single relation, we may retain a file scan (as the cheapest overall plan for fetching all tuples) and a B+ tree index (as the cheapest plan for fetching all tuples in the search key order).

Pass 2: We generate all two-relation plans by considering each single-relation plan retained after Pass 1 as the outer relation and (successively) every other

relation as the inner relation. Suppose that A is the outer relation and B the inner relation for a particular two-relation plan. We examine the list of selections in the WHERE clause and identify:

1. Selections that involve only attributes of B and can be applied before the join.
2. Selections that define the join (i.e., are conditions involving attributes of both A and B and no other relation).
3. Selections that involve attributes of other relations and can be applied only after the join.

The first two groups of selections can be considered while choosing an access path for the inner relation B. We also identify the attributes of B that do not appear in the SELECT clause or in any selection conditions in the second or third group and can therefore be projected out before the join.

Note that our identification of attributes that can be projected out before the join and selections that can be applied before the join is based on the relational algebra equivalences discussed earlier. In particular, we rely on the equivalences that allow us to push selections and projections ahead of joins. As we will see, whether we actually perform these selections and projections ahead of a given join depends on cost considerations. The only selections that are really applied *before* the join are those that match the chosen access paths for A and B. The remaining selections and projections are done on-the-fly as part of the join.

An important point to note is that tuples generated by the outer plan are assumed to be *pipelined* into the join. That is, we avoid having the outer plan write its result to a file that is subsequently read by the join (to obtain outer tuples). For SOme join methods, the join operator might require materializing the outer tuples. For example, a hash join would partition the incoming tuples, and a sort-merge join would sort them if they are not already in the appropriate sort order. Nested loops joins, however, can use outer tuples as they are generated and avoid materializing them. Similarly, sort-merge joins can use outer tuples as they are generated if they are generated in the sorted order required for the join. We include the cost of materializing the outer relation, should this be necessary, in the cost of the join. The adjustments to the join costs discussed in Chapter 14 to reflect the use of pipelining or materialization of the outer are straightforward.

For each single-relation plan for A retained after Pass 1, for each join method that we consider, we must determine the best access method to use for B. The access method chosen for B retrieves, in general, a subset of the tuples in B, possibly with some fields eliminated, as discussed later. Consider relation B.

We have a collection of selections (some of which are the join conditions) and projections on a single relation, and the choice of the best access method is made as per the discussion in Section 15.4.1. The only additional consideration is that the join method might require tuples to be retrieved in some order. For example, in a sort-merge join, we want the inner tuples in sorted order on the join column(s). If a given access method does not retrieve inner tuples in this order, we must add the cost of an additional sorting step to the cost of the access method.

Pass 3: We generate all three-relation plans. We proceed as in Pass 2, except that we now consider plans retained after Pass 2 as outer relations, instead of plans retained after Pass 1.

Additional Passes: This process is repeated with additional passes until we produce plans that contain all the relations in the query. We now have the cheapest overall plan for the query as well as the cheapest plan for producing the answers in some interesting order.

If a multiple-relation query contains a GROUP BY clause and aggregate functions such as MIN, MAX, and SUM in the SELECT clause, these are dealt with at the very end. If the query block includes a GROUP BY clause, a set of tuples is computed based on the rest of the query, as described above, and this set is sorted as per the GROUP BY clause. Of course, if there is a plan according to which the set of tuples is produced in the desired order, the cost of this plan is compared with the cost of the cheapest plan (assuming that the two are different) plus the sorting cost. Given the sorted set of tuples, partitions are identified and any aggregate functions in the SELECT clause are applied on a per-partition basis, as per the discussion in Chapter 14.

Examples of Multiple-Relation Query Optimization

Consider the query tree shown in Figure 12.3. Figure 15.6 shows the same query, taking into account how selections and projections are considered early.

In looking at this figure, it is worth emphasizing that the selections shown on the leaves are not necessarily done in a distinct step that precedes the join—rather, as we have seen, they are considered as potential matching predicates when considering the available access paths on the relations.

Suppose that we have the following indexes, all unclustered and using Alternative (2) for data entries: a B+ tree index on the *rating* field of Sailors, a hash index on the *sid* field of Sailors, and a B+ tree index on the *bid* field of

Optimization in Commercial Systems: IBM DB2, Informix, Microsoft SQL Server, Oracle 8, and Sybase ASE all search for left-deep trees using dynamic programming, as described here, with several variations. For example, Oracle always considers interchanging the two relations in a hash join, which could lead to right-deep trees or hybrids. DB2 generates some bushy trees as well. Systems often use a variety of strategies for generating plans, going beyond the systematic bottom-up enumeration that we described, in conjunction with a dynamic programming strategy for costing plans and remembering interesting plans (to avoid repeated analysis of the same plan). Systems also vary in the degree of control they give users. Sybase ASE and Oracle 8 allow users to force the choice of join orders and indexes--Sybase ASE even allows users to explicitly edit the execution plan-whereas IBM DB2 does not allow users to direct the optimizer other than by setting an 'optimization level,' which influences how many alternative plans the optimizer considers.

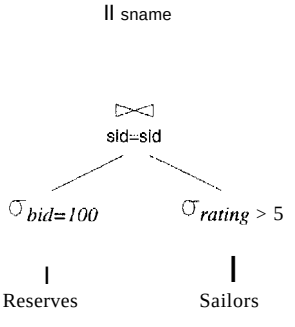


Figure 15.6 A Query Tree

Reserves. In addition, we assume that we can do a sequential scan of both Reserves and Sailors. Let us consider how the optimizer proceeds.

In Pass 1, we consider three access methods for Sailors (B+ tree, hash index, and sequential scan), taking into account the selection $\sigma_{rating>5}$. This selection matches the B+ tree on *rating* and therefore reduces the cost for retrieving tuples that satisfy this selection. The cost of retrieving tuples using the hash index and the sequential scan is likely to be much higher than the cost of using the B+ tree. So the plan retained for Sailors is access via the B+ tree index, and it retrieves tuples in sorted order by *rating.* Similarly, we consider two access methods for Reserves taking into account the selection $\sigma_{bid=100}$. This selection matches the B+ tree index on Reserves, and the cost of retrieving matching tuples via this index is likely to be much lower than the cost of retrieving tuples using a sequential scan; access through the B+ tree index is therefore the only plan retained for Reserves after Pass 1.

In Pass 2, we consider taking the (relation computed by the) plan for Reserves and joining it (as the outer) with Sailors. In doing so, we recognize that now, we need only Sailors tuples that satisfy $\sigma_{rating>5}$ and $\sigma_{sid=value}$, where *value* is some value from an outer tuple. The selection $\sigma_{sid=value}$ matches the hash index on the *sid* field of Sailors, and the selection $\sigma_{rating>5}$ matches the B+ tree index on the *rating* field. Since the equality selection has a much lower reduction factor, the hash index is likely to be the cheaper access method. In addition to the preceding consideration of alternative access methods, we consider alternative join methods. All available join methods are considered. For example, consider a sort-merge join. The inputs must be sorted by *sid;* since neither input is sorted by *sid* or has an access method that can return tuples in this order, the cost of the sort-merge join in this case must include the cost of storing the two inputs in temporary relations and sorting them. A sort-merge join provides results in sorted order by *sid,* but this is not a useful ordering in this example because the projection π_{sname} is applied (on-the-fly) to the result of the join, thereby eliminating the *sid* field from the answer. Therefore, the plan using sort-merge join is retained after Pass 2 only if it is the least expensive plan involving Reserves and Sailors.

Similarly, we also consider taking the plan for Sailors retained after Pass 1 and joining it (as the outer relation) with Reserves. Now we recognize that we need only Reserves tuples that satisfy $\sigma_{bid=100}$ and $\sigma_{sid=value}$, where *value* is some value from an outer tuple. Again, we consider all available join methods.

We finally retain the cheapest plan overall.

As another example, illustrating the case when more than two relations are joined, consider the following query:

```
SELECT    S.sid, COUNT(*) AS numres
FROM      Boats B, Reserves R, Sailors S
WHERE     R.sid = S.sid AND B.bid=R.bid AND Rcolor = 'red'
GROUP BY S.sid
```

This query finds the number of red boats reserved by each sailor. This query is shown in the form of a tree in Figure 15.7.

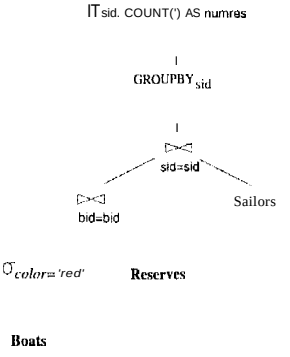


Figure 15.7 A Query Tree

Suppose that the following indexes are available: for Reserves, a B+ tree on the *sid* field and a clustered B+ tree on the *bid* field; for Sailors, a B+ tree index on the *sid* field and a hash index on the *sid* field; and for Boats, a B+ tree index on the *color* field and a hash index on the *color* field. (The list of available indexes is contrived to create a relatively simple, illustrative example.) Let us consider how this query is optimized. The initial focus is on the SELECT, FROM, and WHERE clauses.

In Pass 1, the best plan is found for accessing each relation, regarded as the first relation in an execution plan. :For Reserves and Sailors, the best plan is obviously a. file scan because no selections match an available index. The best plan for Boats is to use the hash index on *color,* which matches the selection *B. coloT = 'red'.* The B+ tree on *color* also matches this selection and is retained even though the hash index is cheaper, because it returns tuples in sorted order by *color.*

In Pass 2, for each of the plans generated in Pass 1, taken as the outer relation, we consider joining another rela.tion as the inner one. Hence, we consider each of the following joins: file scan of Reserves (outer) with Boats (inner), file scan of lleserves (outer) with Sailors (inner), file scan of Sailors (outer) with Boats (inner), file scan of Sailors (outer) with Reserves (inner), Boats accessed via B+ tree index on *color* (outer) with Sailors (inner), Boats accessed via hash

index on *color* (outer) with Sailors (inner), Boats accessed via B+ tree index on *color* (outer) with Reserves (inner), and Boats accessed via hash index on *color* (outer) with Reserves (inner).

For each such pair, we consider every join method, and for each join method, we consider every available access path for the inner relation. For each pair of relations, we retain the cheapest of the plans considered for every sorted order in which the tuples are generated. For example, with Boats accessed via the hash index on *coloT* as the outer relation, an index nested loops join accessing Reserves via the B+ tree index on *bid* is likely to be a good plan; observe that there is no hash index on this field of Reserves. Another plan for joining Reserves and Boats is to access Boats using the hash index on *coloT*, access Reserves using the B+ tree on *bid*, and use a sort-merge join; this plan, in contrast to the previous one, generates tuples in sorted order by *bid*. It is retained even if the previous plan is cheaper, unless an even cheaper plan produces the tuples in sorted order by *bid*. However, the previous plan, which produces tuples in no particular order, would not be retained if this plan is cheaper.

A good heuristic is to avoid considering cross-products if possible. If we apply this heuristic, we would not consider the following 'joins' in Pass 2 of this example: file scan of Sailors (outer) with Boats (inner), Boats accessed via B+ tree index on *color* (outer) with Sailors (inner), and Boats accessed via hash index on *color* (outer) with Sailors (inner).

In Pass 3, for each plan retained in Pass 2, taken as the outer relation, we consider how to join the remaining relation as the inner one. An example of a plan generated at this step is the following: Access Boats via the hash index on *coloT*, access Reserves via the B+ tree index on *bid*, and join them using a sort-merge join, then take the result of this join as the outer and join with Sailors using a sort-merge join, accessing Sailors via the B+ tree index on the *sid* field. Note that, since the result of the first join is produced in sorted order by *bid*, whereas the second join requires its inputs to be sorted by *s'id*, the result of the first join must be sorted by *sid* before being used in the second join. The tuples in the result of the second join are generated in sorted order by *sid*.

The GROUP BY clause is considered after all joins, and it requires sorting on the *sid* field. For each plan retained in Pass 3, if the result is not sorted on *sid*, we add the cost of sorting on the *sid* field. The sample plan generated in Pass 3 produces tuples in *sid* order; therefore, it may be the cheapest plan for the query even if a cheaper plan joins all three relations but does not produce tuples in *sid* order.

15.5 NESTED SUBQUERIES

The unit of optimization in a typical system is a *query block,* and nested queries are dealt with using some form of nested loops evaluation. Consider the following nested query in SQL: *Find the names of sailors with the highest rating:*

```
SELECT S.sname
FROM   Sailors S
WHERE  S.rating = ( SELECT MAX (S2.rating)
                    FROM   Sailors S2 )
```

In this simple query, the nested subquery can be evaluated just once, yielding a single value. This value is incorporated into the top-level query as if it had been part of the original statement of the query. For example, if the highest rated sailor has a rating of 8, the WHERE clause is effectively modified to WHERE *S.rating* = 8.

However, the subquery sometimes returns a relation, or more precisely, a table in the SQL sense (i.e., possibly with duplicate rows). Consider the following query: *Find the names of sailors who have reserved boat number 103:*

```
SELECT S.sname
FROM   Sailors S
WHERE  S.sid IN ( SELECT Rsid
                  FROM   Reserves R
                  WHERE  Rbid = 103 )
```

Again, the nested subquery can be evaluated just once, yielding a collection of *sids.* For each tuple of Sailors, we must now check whether the *sid* value is in the computed collection of *sids;* this check entails a join of Sailors and the computed collection of *sids,* and in principle we have the full range of join methods to choose from. For example, if there is an index on the *sid* field of Sailors, an index nested loops join with the computed collection of *sids* as the outer relation and Sailors as the inner one might be the most efficient join method. However, in many systems, the query optimizer is not smart enough to find this strategy a common approach is to always do a nested loops join in which the inner relation is the collection of *sids* computed from the subquery (and this collection may not be indexed).

The motivation for this approach is that it is a simple variant of the technique used to deal with *correlated queries* such as the following version of the previous query:

```
SELECT S.snallle
```

```
FROM    Sailors S
WHERE   EXISTS ( SELECT *
                 FROM   Reserves R
                 WHERE  R.bid = 103
                        AND S.sid = R.sid )
```

This query is *correlated-"the* tuple variable *S* from the top-level query appears in the nested subquery. Therefore, we cannot evaluate the subquery just once. In this case the typical evaluation strategy is to evaluate the nested subquery for each tuple of Sailors.

An important point to note about nested queries is that a typical optimizer is likely to do a poor job, because of the limited approach to nested query optimization. This is highlighted next:

- In a nested query with correlation, the join method is effectively index nested loops, with the inner relation typically a subquery (and therefore potentially expensive to compute). This approach creates two distinct problems. First, the nested subquery is evaluated once per outer tuple; if the same value appears in the correlation field *(S.sid* in our example) of several outer tuples, the same subquery is evaluated many times. The second problem is that the approach to nested subqueries is not *set-oriented.* In effect, a join is seen as a scan of the outer relation with a selection on the inner subquery for each outer tuple. This precludes consideration of alternative join methods, such as a sort-merge join or a hash join, that could lead to superior plans.
- Even if index nested loops is the appropriate join method, nested query evaluation may be inefficient. For example, if there is an index on the *sid* field of Reserves, a good strategy might be to do an index nested loops join with Sailors as the outer relation and Reserves as the inner relation and apply the selection on *bid* on-the-fly. However, this option is not considered when optimizing the version of the query that uses IN, because the nested subquery is fully evaluated as a first step; that is, Reserves tuples that meet the *bid* selection are retrieved first.
- Opportunities for finding a good evaluation plan may also be missed because of the implicit ordering imposed by the nesting. For example, if there is an index. on the *sid* field of Sailors, an index nested loops join with Reserves as the outer relation and Sailors as the inner one might be the most efficient plan for our example correla,ted query. However. this join ordering is never considered by an optimizer.

A nested query often has an equivalent query without nesting, and a correlated query often has an equivalent query without correlation. We already saw cor-

Nested Queries: IBM DB2, Informix, Microsoft SQL Server, Orade 8, and Sybase ASE all use some version of correlated evaluation to handle nested queries, which are an important part of the TPC-D benchmark; IBM and Informix support a version in which the results of subqueries are stored in a 'memo' table and the same subquery is not executed multiple times. All these RDBMSs consider decqrrelation and "flattening" of nested queries as an option. Microsoft SQL Server, Oracle 8 and IBM DB2 also use rewriting techniques, e.g., Magic Sets (see Chapter 24) or variants, in conjunction with decorrelation.

related and uncorrelated versions of the example nested query. There is also an equivalent query without nesting:

```
SELECT  S.sname
FROM    Sailors S, Reserves R
WHERE   S.sid = R.sid AND R.bid=103
```

A typical SQL optimizer is likely to find a much better evaluation strategy if it is given the unnested or 'decOlTelated' version of the example query than if it were given either of the nested versions of the query. Many current optimizers cannot recognize the equivalence of these queries and transform one of the nested versions to the nonnested form. This is, unfortunately, up to the educated user. From an efficiency standpoint, users are advised to consider such alternative formulations of a query.

We conclude our discussion of nested queries by observing that there could be several levels of nesting. In general, the approach we sketched is extended by evaluating such queries from the innermost to the outermost levels, in order, in the absence of correlation. A correlated subquery must be evaluated for each candidate tuple of the higher-level (sub)query that refers to it. The basic idea is therefore similar to the case of one-level nested queries; we omit the details.

15.6 THE SYSTEM R OPTIMIZER

Current relational query optimizers have been greatly influenced by choices made in the design of IBM's System R query optimizer. Important design choices in the System R optimizer include:

1. The use of *statistics* about the database instance to estiInate the cost of a query evaluation plan.
2. A decision to consider only plans with binary joins in which the inner relation is a base relation (i.e., not a telnporary relation). This heuristic

reduces the (potentially very large) number of alternative plans that must be considered.

3. A decision to focus optimization on the class of SQL queries without nesting and treat nested queries in a relatively ad hoc way.
4. A decision not to perform duplicate elimination for projections (except as a final step in the query evaluation when required by a DISTINCT clause).
5. A model of cost that accounted for CPU costs as well as I/O costs.

Our discussion of optimization reflects these design choices, except for the last point in the preceding list, which we ignore to retain our simple cost model based on the number of page l/Os.

15.7 OTHER APPROACHES TO QUERY OPTIMIZATION

We have described query optimization based on an exhaustive search of a large space of plans for a given query. The space of all possible plans grows rapidly with the size of the query expression, in particular with respect to the number of joins, because join-order optimization is a central issue. Therefore, heuristics are used to limit the space of plans considered by an optimizer. A widely used heuristic is that only left-deep plans are considered, which works well for most queries. However, once the number of joins becomes greater than about 15, the cost of optimization using this exhaustive approach becomes prohibitively high, even if we consider only left-deep plans.

Such complex queries are becoming important in decision-support environments, and other approaches to query optimization have been proposed. These include rule-based optimizers, which use a set of rules to guide the generation of candidate plans, and randomized plan generation, which uses probabilistic algorithms such as *simulated annealing* to explore a large space of plans quickly, with a reasonable likelihood of finding a good plan.

Current research in this area also involves techniques for estimating the size of intermediate relations more accurately; parametric query optimization, which seeks to find good plans for a given query for each of several different conditions that might be encountered at run-time; and multiple-query optimization, in which the optimizer takes concurrent execution of several queries into account.

15.8 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- \What is an *SQL query block*? Why is it important in the context of query optimization? (Section 15.1)
- Describe how a query block is translated into extended relational algebra. Describe and motivate the extensions to relational algebra. Why are $\sigma\pi$x expressions the focus of an optimizer? (Section 15.1)
- \What are the two parts to estimating the cost of a query plan? (Section 15.2)
- How is the result size estimated for a $\sigma\pi\times$ expression? Describe the use of *reduction factors,* and explain how they are calculated for different kinds of selections? (Section 15.2.1)
- What are *histograms?* How do they help in cost estimation? Explain the differences between the different kinds of histograms, with particular attention to the role of frequent data values. (Section 15.2.1)
- When are two relational algebra expressions considered *equivalent?* How is equivalence used in query optimization? What algebra equivalences that justify the common optimizations of pushing selections ahead of joins and re-ordering join expressions? (Section 15.3)
- Describe *left-deep* plans and explain why optimizers typically consider only such plans. (Section 15.4)
- What plans are considered for (sub)queries with a single relation? Of these, which plans are retained in the dynamic programming approach to enumerating left-deep plans? Discuss access methods and output order in your answer. In particular, explain *index-only plans* and why they are attractive. (Section 15.4)
- Explain how query plans are generated for queries with multiple relations. Discuss the space and time complexity of the dynamic programming approach, and how the plan generation process incorporates heuristics like pushing selections and join ordering. How are index-only plans for multiple-relation queries identified? How are pipelining opportunities identified? (Section 15.4)
- How are nested subqueries optimized and evaluated? Discuss correlated queries and the additional optimization challenges they present. \Vhy are plans produced for nested queries typically of poor quality? What is the lesson for application programmers? (Section 15.5)
- Discuss some of the influential design choices made in the System R optimizer. (Section 15.6)
- Briefly survey optimization techniques that go beyond the dynamic programming framework discussed in this chapter. (Section 15.7)

EXERCISES

Exercise **15.1** Briefly answer the following questions:

1. In the context of query optimization, what is an *SQL query block?*
2. Define the term *reduction factor.*
3. Describe a situation in which projection should precede selection in processing a project-select query, and describe a situation where the opposite processing order is better. (Assume that duplicate elimination for projection is done via sorting.)
4. If there are unclustered (secondary) B+ tree indexes on both *R.a* and *S.b,* the join $R \bowtie_{a=b} S$ could be processed by doing a sort-merge type of join-without doing any sorting-by using these indexes.
   (a) Would this be a good idea if Rand S each has only one tuple per page or would it be better to ignore the indexes and sort Rand S? Explain.
   (b) What if Rand S each have many tuples per page? Again, explain.
5. Explain the role of *interesting orders* in the System R optimizer.

Exercise **15.2** Consider a relation with this schema:

Ernployees(eid: integer, *ename:* string, *sal:* integer, *title:* string, *age:* integer)

Suppose that the following indexes, all using Alternative (2) for data entries, exist: a hash index on *eid,* a B+ tree index on *sal,* a hash index on *age,* and a clustered B+ tree index on *(age, sal).* Each Employees record is 100 bytes long, and you can assume that each index data entry is 20 bytes long. The Employees relation contains 10,000 pages.

1. Consider each of the following selection conditions and, assuming that the reduction factor (RF) for each term that matches an index is 0.1, compute the cost of the most selective access path for retrieving all Employees tuples that satisfy the condition:
   (a) $sol > 100$
   (b) $age = 25$
   (c) $age > 20$
   (d) $eid = 1,000$
   (e) $sal > 200 \wedge age > 30$
   (f) $sal > 200 \wedge age = 20$
   (g) $sal > 200 \wedge title = 'CFO'$
   (h) $sal > 200 \wedge age > 30 \wedge title = 'CFO'$
2. Suppose that, for each of the preceding selection conditions, you want to retrieve the average salary of qualifying tuples. For each selection condition, describe the least expensive evaluation method and state its cost.
3. Suppose that, for each of the preceding selection conditions, you want to compute the average salary for each *age* group. For each selection condition, describe the least expensive evaluation method and state its cost.
4. Suppose that, for each of the preceding selection conditions, you want to compute the average age for each sa/level (Le.) group by *sal).* For each selection condition, describe the least expensive evaluation method and state its cost.

5. For each of the following selection conditions, describe the best evaluation method:
   (a) $sal > 200 \vee age = 20$
   (b) $sal > 200 \vee title = 'CFO'$
   (c) $title = 'CFO' \wedge ename = 'Joe'$

Exercise **15.3** For each of the following SQL queries, for each relation involved, list the attributes that must be examined to compute the answer. All queries refer to the following relations:

Emp(eid: integer, *did:* integer, *sal:* integer, *hobby:* char(20))
Dept(*did:* integer, *dname:* char(20), *floor:* integer, *budget:* real)

1. SELECT COUNT(*) FROM Emp E, Dept D WHERE E.did = D.did
2. SELECT MAX(E.sal) FROM Emp E, Dept D WHERE E.did = D.did
3. SELECT MAX(E.sal) FROM Emp E, Dept D WHERE E.did = D.did AND D.floor = 5
4. SELECT E.did, COUNT(*) FROM Emp E, Dept D WHERE E.did = D.did GROUP BY D.did
5. SELECT D.floor, AVG(D.budget) FROM Dept D GROUP BY D.tloor HAVING COUNT(*) > 2
6. SELECT D.tloor, AVG(D.budget) FROM Dept D GROUP BY D.floor ORDER BY D.floor

Exercise **15.4** You are given the following information:

Executives has attributes *ename, title, dname,* and *address;* all are string fields of the same length.
The *ename* attribute is a candidate key.
The relation contains 10,000 pages.
There are 10 buffer pages.

1. Consider the following query:

   SELECT E.title, E.ename FROM Executives E WHERE E.title='CFO'

   Assume that only 10% of Executives tuples meet the selection condition.

   (a) Suppose that a clustered B+ tree index on *title* is (the only index) available. What is the cost of the best plan? (In this and subsequent questions, be sure to describe the plan you have in mind.)
   (b) Suppose that an unclustered B+ tree index on *title* is (the only index) available. What is the cost of the best plan?
   (c) Suppose that a clustered B+ tree index on *enarne* is (the only index) available. What is the cost of the best plan?
   (d) Suppose that a clustered B+ tree index on *address* is (the only index) available. What is the cost of the best pian?
   (e) Suppose that a clustered B+ tree index on *(ename, title)* is (the only index) available. What is the cost of the best plan?

2. Suppose that the query is as follows:

   SELECT E.ename FROM Executives E WHERE E.title='CFO' AND E.dname='Toy'

Assume that only 10% of Executives tuples IIleet the condition $E.title = 'CFO'$, only 10% meet $E.dname = 'Toy'$, and that only 5% meet both conditions.

(a) Suppose that a clustered B+ tree index on *title* is (the only index) available. What is the cost of the best plan?

(b) Suppose that a clustered B+ tree index on *dname* is (the only index) available. What is the cost of the best plan?

(c) Suppose that a clustered B+ tree index on *(title, dname)* is (the only index) available. What is the cost of the best plan?

(d) Suppose that a clustered B+ tree index on *(title, ename)* is (the only index) available. What is the cost of the best plan?

(e) Suppose that a clustered B+ tree index on *(dname, title, ename)* is(the only index) available. What is the cost of the best plan?

(f) Suppose that a clustered B+ tree index on *(ename, title, dname)* is (the only index) available. What is the cost of the best plan?

3. Suppose that the query is as follows:

```
SELECT E.title, COUNT(*) FROM Executives E GROUP BY E.title
```

(a) Suppose that a clustered B+ tree index on *title* is (the only index) available. What is the cost of the best plan?

(b) Suppose that an unclustered B+ tree index on *t'ltle* is (the only index) available. What is the cost of the best plan?

(c) Suppose that a clustered B+ tree index on *ename* is (the only index) available. What is the cost of the best plan?

(d) Suppose that a clustered B+ tree index on *(ename, title)* is (the only index) available. What is the cost of the best plan?

(e) Suppose that a clustered B+ tree index on *(title, ename)* is (the only index) available. What is the cost of the best plan?

4. Suppose that the query is as follows:

```
SELECT E.title, COUNT(*) FROM Executives E
WHERE E.dname > 'W%' GROUP BY E.title
```

Assume that only 10% of Executives tuples meet the selection condition.

(a) Suppose that a clustered B+ tree index on *title* is (the only index) available. What is the cost of the best plan? If an additional index (on any search key you want) is available, would it help produce a better plan?

(b) Suppose that an unclustered B+ tree index on *title* is (the only index) available. What is the cost of the best plan?

(c) Suppose. that a clustered B+ tree index on *dname* is (the only index) available. What is the cost of the best plan? If an additional index (on any search key you want) is available, would it help to produce a better plan'?

(d) Suppose that a clustered B+ tree index on *(dname, title)* is (the only index) available. What is the cost of the best plan?

(e) Suppose that a clustered B+ tree index on *(title,dname)* is (the only index) available. What is the cost of the best plan?

Exercise **15.5** Consider the query $\pi_{A,B,C,D}(R \bowtie_{A=C} S)$. Suppose that the projection routine is based on sorting and is smart enough to eliminate all but the desired attributes during the initial pass of the sort and also to toss out duplicate tuples on-the-fly while sorting, thus eliminating two potential extra passes. Finally, assume that you know the following:

R is 10 pages long, and R tuples are 300 bytes long.
S is 100 pages long, and S tuples are 500 bytes long.
C is a key for S, and A is a key for R.
The page size is 1024 bytes.
Each S tuple joins with exactly one R tuple.
The combined size of attributes A, B, G, and D is 450 bytes.
A and B are in R and have a combined size of 200 bytes; C and D are in S.

1. What is the cost of writing out the final result? (As usual, you should ignore this cost in answering subsequent questions.)
2. Suppose that three buffer pages are available, and the only join method that is implemented is simple (page-oriented) nested loops.
   (a) Compute the cost of doing the projection followed by the join.
   (b) Compute the cost of doing the join followed by the projection.
   (c) Compute the cost of doing the join first and then the projection on-the-fly.
   (d) Would your answers change if 11 buffer pages were available?

Exercise 15.6 Briefly answer the following questions:

1. Explain the role of relational algebra equivalences in the System R optimizer.
2. Consider a relational algebra expression of the form $\sigma_c(\pi_l(R \times S))$. Suppose that the equivalent expression with selections and projections pushed as much as possible, taking into accollnt only relational algebra equivalences, is in one of the following forms. In each case give an illustrative example of the selection conditions and the projection lists (c, l, el, 11, etc.).
   (a) *Equivalent maximally pushed form:* $\pi_{l1}(\sigma_{c1}(R) \times S)$,
   (b) *Equivalent maximally pushed form:* $\pi_{l1}(\sigma_{c1}(R) \times \sigma_{c2}(S))$.
   (c) *Equivalent maximally pushed form:* $\sigma_c(\pi_{l1}(\pi_{l2}(R) \times 8))$.
   (d) *Equivalent maximally pushed fONT!:* $\sigma_{c1}(\pi_{l1}(\sigma_{c2}(\pi_{l2}(R)) \times 8))$.
   (e) *Equivalent ma:rimally pushed form:* $\sigma_{c1}(\pi_{l1}(\pi_{l2}(\sigma_{c2}(R)) \times S))$.
   (f) *Equivalent maximally pushed form:* $\pi_l(\sigma_{c1}(\pi_{l1}(\pi_{l2}(\sigma_{c2}(R)) \times 8)))$.

Exercise 15.7 Consider the following relational schema and SQL query. The schema captures information about employees, departments, and company finances (organized on a per department basis).

Emp(eid: integer, *did:* integer, *sal:* integer, *hobby:* char(20))
Dept(did: integer, *dname:* char(20), *floor:* integer, *phone:* char(10))
Finance(*did:* integer, *budget:* real, *sales:* real, *expenses:* real)

Consider the following query:

```
SELECT D.dname, F.budget
FROM   Emp E, Dept D, Finance F
WHERE  E.did=D.did AND D.did=F.did AND D.floor=1
       AND E.sal ≥ 59000 AND E.hobby = 'yodeling'
```

1. Identify a relational algebra tree (or a relational algebra expression if you prefer) that reflects the order of operations a decent query optimizer would choose.
2. List the join orders (i.e., orders in which pairs of relations can be joined to compute the query result) that a relational query optimizer will consider. (Assume that the optimizer follows the heuristic of never considering plans that require the computation of cross-products.) Briefly explain how you arrived at your list.
3. Suppose that the following acldditional information is available: lJnclustered B+ tree indexes exist on *Emp.did*, *Ernp.sal*, *Dept.floor*, *Dept.did*, and *Finance.did*. The system's statistics indicate that employee salaries range from 10,000 to 60,000, employees enjoy 200 different hobbies, and the company owns two floors in the building. There are a total of 50,000 employees and 5,000 departments (each with corresponding financial information) in the database. The DBMS used by the company has just one join method available, index nested loops.
   (a) For each of the query's base relations (Emp, Dept, and Finance) estimate the number of tuples that would be initially selected from that relation if all of the non-join predicates on that relation were applied to it before any join processing begins.
   (b) Given your answer to the preceding question, which of the join orders considered by the optimizer has the lowest estimated cost?

Exercise 15.8 Consider the following relational schema and SQL query:

Suppliers(sid: integer, *snarne:* char(20), *city:* char(20»)
Supply(*sid:* integer, *pid:* integer)
Parts(pid: integer, *pnarne:* char(20), *price:* real)

```
SELECT S.sname, P.pname
FROM   Suppliers S, Parts P, Supply Y
WHERE  S.sid = Y.sid AND Y.pid = P.pid AND
       S.city = 'Madison' AND P.price ≤ 1,000
```

1. What information abollt these relations does the query optimizer need to select a good query execution plan for the given query?
2. How many different join orders, assuming that cross-products are disallowed, does a System R. style query optimizer consider whcn deciding how to process the given query? List each of thcse join orders.
3. What indexes might be of help in processing this query? Explain briefly.
4. How does adding DISTINCT to the SELECT clause affect the plans produced?
5. How does adding ORDER BY *sname* to the query affect the plans produced?
6. How does adding GROUP BY *sname* to the query affect the plans produced?

Exercise 15.9 Consider the following scenario:

Emp(*eid:* integer, *sal:* integer, *age:* real, *did:* integer)
Dept(*did:* integer, *projid:* integer, *budget:* real, *status:* char(10))
Proj(*projid:* integer, *code:* integer, *report:* varchar)

Assume that each Emp record is 20 bytes long, each Dept record is 40 bytes long, and each Proj record is 2000 bytes long on average. There are 20,000 tuples in Emp, 5000 tuples in Dept (note that *did* is not a key), and 1000 tuples in Proj. Each department, identified by *did,* has 10 projects on average. The file system supports 4000 byte pages, and 12 buffer pages are available. All following questions are based on this information. You can assume uniform distribution of values. State any additional assumptions. The cost metric to use is *the number of page I/Os.* Ignore the cost of writing out the final result.

1. Consider the following two queries: "Find all employees with *age* = 30" and "Find all projects with *code* = 20," Assume that the number of qualifying tuples is the same in each case. If you are building indexes on the selected attributes to speed up these queries, for which query is a *clustered* index (in comparison to an *unclustered* index) more important?
2. Consider the following query: "Find all employees with *age* > 30." Assume that there is an unclustered index on *age.* Let the number of qualifying tuples be *N.* For what values of *N* is a sequential scan cheaper than using the index?
3. Consider the following query:

```
SELECT *
FROM   Emp E, Dept D
WHERE  E.did=D.did
```

   (a) Suppose that there is a clustered hash index on *did* on Emp. List all the plans that are considered and identify the plan with the lowest estimated cost.

   (b) Assume that both relations are sorted on the join column. List all the plans that are considered and show the plan with the lowest estimated cost.

   (c) Suppose that there is a clustered B+ tree index on *did* on Emp and Dept is sorted on *did.* List all the plans that are considered and identify the plan with the lowest estimated cost.

4. Consider the following query:

```
SELECT      D.dicl, COUNT(*)
FROM        Dept D, Proj P
WHERE       D.projid=P.projid
GROUP BY    D.clid
```

   (a) Suppose that no indexes are available. Show the plan with the lowest estimated cost.

   (b) If there is a hash index on *P.projid* what is the plan with lowest estimated cost?

   (c) If there is a hash index on *D.projid* what is the plan with lowest estimated cost?

   (d) If there is a hash index on *D-JiTojid* and *P.projid* what is the plan with lowest estimated cost?

   (e) Suppose that there is a clustered B+ tree index on *D.did* and a hash index on *P.]Jmjid.* Show the plan with the lowest estimated cost.

   (f) Suppose that there is a clustered B+ tree index on *D.did,* a hash index on *D.]JT'O)id,* and a hash index on *P.projid.* Show the plan with the lowest estimated cost.

(g) Suppose that there is a clustered B+ tree index on *(D.did, D.projid)* and a hash index on *P.pmjid.* Show the plan with the lowest estimated cost.

(h) Suppose that there is a clustered B+ tree index on *⟨D.projid, D.did⟩* and a hash index on *P.pmjid.* Show the plan with the lowest estimated cost.

5. Consider the following query:

```
SELECT      D.did, COUNT(*)
FROM        Dept D, Proj P
WHERE       D.projid=P.projid AND D.budget>99000
GROUP BY    D.did
```

Assume that department budgets are uniformly distributed in the range 0 to 100,000.

(a) Show the plan with lowest estimated cost if no indexes are available.

(b) If there is a hash index on *P.pmjid* show the plan with lowest estimated cost.

(c) If there is a hash index on *D. budget* show the plan with lowest estimated cost.

(d) If there is a hash index on *D.pmjid* and *D.budget* show the plan with lowest estimated cost.

(e) Suppose that there is a clustered B+ tree index on *(D.did,D.budget)* and a hash index on *P.projid.* Show the plan with the lowest estimated cost.

(f) Suppose there is a clustered B+ tree index on *D.did,* a hash index on *D.b1ldget,* and a hash index on *P.projid.* Show the plan with the lowest estimated cost.

(g) Suppose there is a clustered B+ tree index on *(D. did, D.budgct, D.projid⟩* and a hash index on *P.pmjid.* Show the plan with the lowest estimated cost.

(h) Suppose there is a clustered B+ tree index on *(D. did, D.projid, D.budget)* and a hash index on *P.pmjid.* Show the plan with the lowest estimated cost.

6. Consider the following query:

```
SELECT  E.eid, D.did, P.projid
FROM    Emp E, Dept D, Proj P
WHERE   E.sal=50,000 AND D.budget>20,000
        E.did=D.did AND D.projid=P.projid
```

Assume that employee salaries are uniformly distributed in the range 10,009 to 110,008 and that project budgets are uniformly distributed in the range 10,000 to 30,000. There is a clustered index on *sal* for Emp, a clustered index on *did* for Dept, and a clustered index on *pmjid* for Proj.

(a) List all the one-relation, two-relation, and three-relation subplans considered in optimizing this query.

(b) Show the plan with the lowest estimated cost for this query.

(c) If the index on Proj were unclustered, would the cost of the preceding plan change substant:ially? What if the index on Emp or on Dept were uncllstered?

BIBLIOGRAPHIC NOTES

Query optimization is critical in a relational DBMS, and it has therefore been extensively studied. We concentrate in this chapter on the approach taken in System R, as described in [668], although our discussion incorporates subsequent refinements to the approach. [784] describes query optimization in Ingres. Good surveys can be found in [410J and [399J. [434] contains several articles on query processing and optimization.

From a theoretical standpoint, [155] shows that determining whether two *conjunctive queries* (queries involving only selections, projections, and cross-products) are equivalent is an NP-complete problem; if relations are *multisets*, rather than sets of tuples, it is not known whether the problem is decidable, although it is Π_2^p hard. The equivalence problem is shown to be decidable for queries involving selections, projections, cross-products, and unions in [643]; surprisingly, this problem is undecidable if relations are multisets [404]. Equivalence of conjunctive queries in the presence of integrity constraints is studied in [30], and equivalence of conjunctive queries with inequality selections is studied in [440].

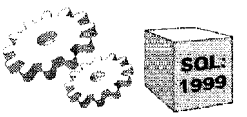
An important problem in query optimization is estimating the size of the result of a query expression. Approaches based on sampling are explored in [352, 353, 384, 481, 569]. The use of detailed statistics, in the form of histograms, to estimate size is studied in [405, 558, 598]. Unless care is exercised, errors in size estimation can quickly propagate and make cost estimates worthless for expressions with several operators, This problem is examined in [400]. [512] surveys several techniques for estimating result sizes and correlations between values in relations. There are a number of other papers in this area; for example, [26, 170, 594, 725], and our list is far from complete,

Semantic query optimization is based on transformations that preserve equivalence only when certain integrity constraints hold. The idea was introduced in [437] and developed further in [148,682, 688].

In recent years, there has been increasing interest in complex queries for decision support applications. Optimization of nested SQL queries is discussed in [298, 426, /130, 557, 760]. The use of the Magic Sets technique for optimizing SQL queries is studied in [553, 554, 555, 670, 673]. Rule-based query optimizers are studiecl in [287, 326, 490, 539, 596]. Finding a good join order for queries with a large number of joins is studied in [401, 402, 453, 726]. Optimization of multiple queries for simultaneous execution is considerecl in [585, 633, 669]. Determining query plans at run-time is discussed in [327, 403]. Re-optimization of running queries based on statistics gathered during query execution is considered by Kabra and DeWitt [413]. Probabilistic optimization of queries is proposed in [183, 229].

PART V

TRANSACTION MANAGEMENT



16

OVERVIEW OF TRANSACTION MANAGEMENT

- What four properties of transactions does a DBMS guarantee?
- Why does a DBMS interleave transactions?
- What is the correctness criterion for interleaved execution?
- What kinds of anomalies can interleaving transactions cause?
- How does a DBMS use locks to ensure correct interleavings?
- What is the impact of locking on performance?
- What SQL commands allow programmers to select transaction characteristics and reduce locking overhead?
- How does a DBMS guarantee transaction atomicity and recovery from system crashes?

-- **Key concepts:** ACID properties, atomicity, consistency, isolation, durability; schedules, serializability, recoverability, avoiding cascading aborts; anomalies, dirty reads, unrepeatable reads, lost updates; locking protocols, exclusive and shared locks, Strict Two-Phase Locking; locking performance, thrashing, hot spots; SQL transaction characteristics, savepoints, rollbacks, phantoms, access mode, isolation level; transaction manager, recovery manager, log, system crash, media failure; stealing frames, forcing pages; recovery phases, analysis, redo and undo.

I always say, keep a diary and someday it'11 keep you.

Mae West

In this chapter, we cover the concept of a *transaction*, which is the foundation for concurrent execution and recovery from system failure in a DBMS. A transaction is defined as *any one execution* of a user program in a DBMS and differs from an execution of a program outside the DBMS (e.g., a C program executing on Unix) in important ways. (Executing the same program several times generates several transactions.)

For performance reasons, a DBMS has to interleave the actions of several transactions. (We motivate interleaving of transactions in detail in Section 16.3.1.) However, to give users a simple way to understand the effect of running their programs, the interleaving is done carefully to ensure that the result of a concurrent execution of transactions is nonetheless equivalent (in its effect on the database) to some serial, or one-at-a-time, execution of the same set of transactions, How the DBMS handles concurrent executions is an important aspect of transaction management and the subject of *concurrency control.* A closely related issue is how the DBMS handles partial transactions, or transactions that are interrupted before they run to normal completion, The DBMS ensures that the changes made by such partial transactions are not seen by other transactions. How this is achieved is the subject of *crash r'ecovery.* In this chapter, we provide a broad introduction to concurrency control and crash recovery in a DBMS, The details are developed further in the next two chapters.

In Section 16.1, we discuss four fundamental properties of database transactions and how the DBMS ensures these properties. In Section 16.2, we present an abstract way of describing an interleaved execution of several transactions, called a *schedule.* In Section 16,3, we discuss various problems that can arise due to interleaved execution, \Ve introduce lock-based concurrency control, the most widely used approach, in Section 16.4. We discuss performance issues associated with lock-based concurrency control in Section 16.5. We consider locking and transaction properties in the context of SQL in Section 16.6, Finally, in Section 16.7, we present an overview of how a clatabase system recovers from crashes and what steps are taken during normal execution to support crash recovery.

16.1 THE ACID PROPERTIES

We introduced the concept of database transactions in Section 1.7, To recapitulate briefly, a transaction is an execution of a user program, seen by the DBMS as a series of read and write operations.

A DBMS must ensure four important properties of transactions to maintain data in the face of concurrent access and system failures:

1. Users should be able to regard the execution of each transaction as atomic: Either all actions are carried out or none are. Users should not have to worry about the effect of incomplete transactions (say, when a system crash occurs).

2. Each transaction, run by itself with no concurrent execution of other transactions, lnust preserve the consistency of the database. The DBMS assumes that consistency holds for each transaction. Ensuring this property of a transaction is the responsibility of the user.

3. Users should be able to understand a transaction without considering the effect of other concurrently executing transactions, even if the DBMS interleaves the actions of several transactions for performance reasons. This property is sometimes referred to as isolation: Transactions are isolated, or protected, from the effects of concurrently scheduling other transactions.

4. Once the DBMS informs the user that a transaction has been successfully completed, its effects should persist even if the system crashes before all its changes are reflected on disk. This property is called durability.

The acronym ACID is sometimes used to refer to these four properties of transactions: atomicity, consistency, isolation and durability. We now consider how each of these properties is ensured in a DBMS.

16.1.1 Consistency and Isolation

Users are responsible for ensuring transaction consistency. That is, the user who submits a transaction must ensure that, when run to completion by itself against a 'consistent' database instance, the transaction will leave the databa.,se in a 'consistent' state. For example, the user may (naturally) have the consistency criterion that fund transfers between bank accounts should not change the total amount of money in the accounts. To transfer money from one account to another, a transaction must debit one account, temporarily leaving the database inconsistent in a global sense, even though the new account balance may satisfy any integrity constraints with respect to the range of acceptable account balances. The user's notion of a consistent database is preserved when the second account is credited with the transferred amount. If a faulty transfer program always credits the second account with one dollar less than the alllount debited frOlll the first account, the DBMS cannot be expected to detect inconsistencies due to such errors in the user program's logic.

The isolation property is ensured by guaranteeing that, even though actions of several transactions rnight be interleaved, the net effect is identical to executing all transactions one after the other in sorne serial order. (We discuss

how the DBMS implements this guarantee in Section 16.4.) For example, if two transactions *T1* and *T2* are executed concurrently, the net effect is guaranteed to be equivalent to executing (all of) *T1* followed by executing *T2* or executing *T2* followed by executing T1. (The DBMS provides no guarantees about which of these orders is effectively chosen.) If each transaction maps a consistent database instance to another consistent database instance, executing several transactions one after the other (on a consistent initial database instance) results in a consistent final database instance.

Database consistency is the property that every transaction sees a consistent database instance. Database consistency follows from transaction atomicity, isolation, and transaction consistency. Next, we discuss how atomicity and durability are guaranteed in a DBMS.

16.1.2 Atomicity and Durability

Transactions can be incomplete for three kinds of reasons. First, a transaction can be **aborted,** or terminated unsuccessfully, by the DBMS because some anomaly arises during execution. If a transaction is aborted by the DBMS for some internal reason, it is automatically restarted and executed anew. Second, the system may crash (e.g., because the power supply is interrupted) while one or more transactions are in progress. Third, a transaction may encounter an unexpected situation (for example, read an unexpected data value or be unable to access some disk) and decide to abort (i.e., terminate itself).

Of course, since users think of transactions as being atomic, a transaction that is interrupted in the middle may leave the database in an inconsistent state. Therefore, a DBMS must find a way to remove the effects of partial transactions from the database. That is, it must ensure transaction atomicity: Either all of a transaction's actions are carried out or none are. A DBMS ensures transaction atomicity by *undoing* the actions of incomplete transactions. This means that users can ignore incomplete transactions in thinking about how the database is modified by transactions over time. To be able to do this, the DBMS maintains a record, called the *log*, of all writes to the database. The log is also used to ensure durability: If the system crashes before the changes made by a completed transaction are written to disk, the log is used to remember and restore these changes when the system restarts.

The DBMS component that ensures atomicity and durability, called the *recovery manager*, is discussed further in Section 16.7.

16.2 TRANSACTIONS AND SCHEDULES

A transaction is seen by the DBMS as a series, or *list*, of **actions.** The actions that can be executed by a transaction include **reads** and **writes** of *database objects.* To keep our notation simple, we assume that an object 0 is always read into a program variable that is also named O. *'Ne* can therefore denote the action of a transaction *T* reading an object 0 as *RT(O);* similarly, we can denote writing as $W_T(O)$. When the transaction *T* is clear from the context, we omit the subscript.

In addition to reading and writing, each transaction *must* specify as its final action either **commit** (i.e., complete successfully) or **abort** (i.e., terminate and undo all the actions carried out thus far). *AbortT* denotes the action of *T* aborting, and *CommitT* denotes *T* committing.

We make two important assumptions:

1. Transactions interact with each other *only* via database read and write operations; for example, they are not allowed to exchange messages.
2. A database is a *fiJ;ed* collection of *independent* objects. When objects are added to or deleted from a database or there are relationships between database objects that we want to exploit for performance, some additional issues arise.

If the first assumption is violated, the DBMS has no way to detect or prevent inconsistencies cause by such external interactions between transactions, and it is upto the writer of the application to ensure that the program is well-behaved. We relax the second assumption in Section 16.6.2.

A **schedule** is a list of actions (reading, writing, aborting, or committing) from a set of transactions, and the order in which two actions of a transaction *T* appear in a schedule must be the same as the order in which they appear in *T*. Intuitively, a schedule represents an actual or potential execution sequence. For example, the schedule in Figure 16.1 shows an execution order for actions of two transactions *T1* and *T2*. We move forward in time as we go down from one row to the next. We emphasize that a schedule describes the actions of transactions *as seen by the DBMS.* In addition to these actions, a transaction rnay carry out other actions, such as reading or writing from operating system files, evaluating arithmetic expressions, and so on; however, we assume that these actions do not affect other transactions; that is, the effect of a transaction on another transaction can be understood solely in terms of the cornmon database objects that they read and write.

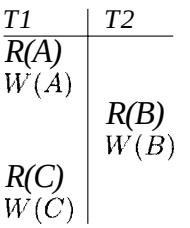


Figure 16.1 A Schedule Involving Two Transactions

Note that the schedule in Figure 16.1 does not contain an abort or commit action for either transaction. A schedule that contains either an abort or a commit for each transaction whose actions are listed in it is called a complete schedule. A complete schedule must contain all the actions of every transaction that appears in it. If the actions of different transactions are not interleaved-that is, transactions are executed from start to finish, one by one-we call the schedule a serial schedule.

16.3 CONCURRENT EXECUTION OF TRANSACTIONS

Now that we have introduced the concept of a schedule, we have a convenient way to describe interleaved executions of transactions. The DBMS interleaves the actions of different transactions to improve performance, but not all interleavings should be allowed. In this section, we consider what interleavings, or schedules, a DBMS should allow.

16.3.1 Motivation for Concurrent Execution

The schedule shown in Figure 16.1 represents an interleaved execution of the two transactions. Ensuring transaction isolation while permitting such concurrent execution is difficult but necessary for performance reasons. First, while one transa.etion is waiting for a page to be read in from disk, the CPU can process another transaction. This is because I/O activity can be done in parallel with CPU activity in a computer. Overlapping I/O and CPU activity reduces the amount of time disks and processors are idle and increases system throughput (the average number of transactions completed in a given time). Second, interleaved execution of a short transaction with a long transaction usually allows the short transaction to complete quickly. In serial execution, a short transaction could get stuck behind a long transaction, leading to unpredictable delays in response time, or average time taken to complete a transaction.

16.3.2 SerializabHity

A **serializable schedule** over a set S of cormnitted transactions is a schedule whose effect on any consistent database instance is guaranteed to be identical to that of some complete serial schedule over S. That is, the database instance that results from executing the given schedule is identical to the database instance that results frOlll executing the transactions in *some* serial order.[1]

As an example, the schedule shown in Figure 16.2 is serializable. Even though the actions of *T1* and *T2* are interleaved, the result of this schedule is equivalent to running *T1* (in its entirety) and then running $T2$. Intuitively, *T1*'s read and write of *B* is not influenced by *T2's* actions on *A*, and the net effect is the same if these actiolls are 'swapped' to obtain the serial schedule *Tl; T2*.

Tl	*T2*
R(A)	
$W(A)$	
	R(A)
	W(A)
R(B)	
$W(B)$	
	R(B)
	W(B)
	Commit
Commit	

Figure 16.2 A Serializable Schedule

Executing transactions serially in different orders may produce different results, but all are presumed to be acceptable: the DBMS makes no guarantees ahout which of them will be the outcome of an interleaved execution. To see this, note that the two example transactions from Figure 16.2 can be interleaved as shown in Figure 16.3. This schedule, also serializable, is equivalent to the serial schedule *T2;* Tl. If *T1* and *T2* are submitted concurrently to a DBMS, either of these schedules (among others) could be chosen.

The preceding definition of a serializable schedule does not cover the case of schedules containing aborted transactions. We extend the definition of serializable schedules to cover aborted transactions in Section 16.3.4.

---

[1] If a transaction prints a value to the screen, this 'effect' is not directly captured in the database. For simplicity, we assume that such values are also written into the database.

T1	T2
	R(A)
	W(A)
R(A)	
	R(B)
	W(B)
W(A)	
R(B)	
W(B)	
	Commit
Commit	

Figure 16.3 Another Serializable Schedule

Finally, we note that a DBMS might sometimes execute transactions in a way that is not equivalent to any serial execution; that is, using a schedule that is not serializable. This can happen for two reasons. First, the DBMS might use a concurrency control method that ensures the executed schedule, though not itself serializable, is equivalent to some serializable schedule (e.g., see Section 17.6.2). Second, SQL gives application programmers the ability to instruct the DBMS to choose non-serializable schedules (see Section 16.6).

16.3.3 Anomalies Due to Interleaved Execution

We now illustrate three main ways in which a schedule involving two consistency preserving, committed transactions could run against a consistent database and leave it in an inconsistent state. Two actions on the same data object conflict if at least one of them is a write. The three anomalous situations can be described in terms of when the actions of two transactions *T1* and *T2* conflict with each other: In a write-read (WR) conflict, *T2* reads a data object previously written by *T1;* we define read-write (RW) and write-write (WW) conflicts similarly.

Reading Uncommitted Data (WR Conflicts)

The first source of anomalies is that a transaction *T2* could read a database object *A* that has been modified by another transaction *T1,* which has not yet committed. Such a read is called a dirty read. A simple example illustrates how such a schedule could lead to an inconsistent database state. Consider two transactions *T1* and *T2*. each of which, run alone, preserves database consistency: *T1* transfers $100 from *A* to *B*, and *T2* increments both *A* and *B* by 6% (e.g., annual interest is deposited into these two accounts). Suppose

that the actions are interleaved so that (1) the account transfer program T1 deducts $100 from account *A*, then (2) the interest deposit program *T2* reads the current values of accounts *A* and *B* and adds 6% interest to each, and then (3) the account transfer program credits $100 to account *B*. The corresponding schedule, which is the view the DBMS has of this series of events, is illustrated in Figure 16.4. The result of this schedule is different from any result that we would get by running one of the two transactions first and then the other. The problem can be traced to the fact that the value of *A* written by *T1* is read by *T2* before *T1* has completed all its changes.

T1	*T2*
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
R(B)	
W(B)	
Commit	

Figure 16.4 Reading Uncommitted Data

The general problem illustrated here is that T1 may write some value into *A* that makes the database inconsistent. As long as *T1* overwrites this value with a 'correct' value of *A* before committing, no harm is done if *T1* and *T2* run in some serial order, because *T2* would then not see the (temporary) inconsistency. On the other hand, interleaved execution can expose this inconsistency and lead to an inconsistent final database state.

Note that although a transaction must leave a database in a consistent state *after* it completes, it is not required to keep the database consistent while it is still in progress. Such a requirement would be too restrictive: To transfer money from one account to another, a transaction *must* debit one account, temporarily leaving the database inconsistent, and then credit the second account, restoring consistency.

Unrepeatable Reads (RW Conflicts)

The second way in which anomalous behavior could result is that a transaction $T2$ could change the value of an object A that has been read by a transaction $T1$, while $T1$ is still in progress.

If $T1$ tries to read the value of A again, it will get a different result, even though it has not modified A in the meantime. This situation could not arise in a serial execution of two transactions; it is called an unrepeatable read.

To see why this can cause problems, consider the following example. Suppose that A is the number of available copies for a book. A transaction that places an order first reads A, checks that it is greater tha,n 0, and then decrements it. Transaction $T1$ reads A and sees the value 1. Transaction $T2$ also reads A and sees the value 1, decrements A to 0 and commits. Transaction $T1$ then tries to decrement A and gets an error (if there is an integrity constraint that prevents A from becoming negative).

This situation can never arise in a serial execution of $T1$ and $T2$; the second transaction would read A and see 0 and therefore not proceed with the order (and so would not attempt to decrement A).

Overwriting Uncommitted Data (WW Conflicts)

The third source of anomalous behavior is that a transaction $T2$ could overwrite the value of an object A, which has already been modified by a transaction $T1$, while $T1$ is still in progress. Even if $T2$ does not read the value of A written by $T1$, a potential problem exists as the following example illustrates.

Suppose that Harry and Larry are two employees, and their salaries must be kept equal. Transaction $T1$ sets their salaries to $2000 and transaction $T2$ sets their salaries to $1000. If we execute these in the serial order $T1$ followed by $T2$, both receive the salary $1000: the serial order $T2$ followed by $T1$ gives each the salary $2000. Either of these is acceptable from a consistency standpoint (although Harry and Larry may prefer a higher salary!). Note that neither transaction reads a salary value before writing it----such a write is called a blind write, tor obvious reasons.

Now, consider the following interleaving of the actions of $T1$ and $T2$: $T2$ sets Harry's salary to $1000, $T1$ sets Larry's salary to $2000, $T2$ sets La.rry's salary to $1000 and commits, and finally $T1$ sets Harry's salary to $2000 and connnits. The result is not identical to the result of either of the two possible serial

executions, and the interleaved schedule is therefore not serializable. It violates the desired consistency criterion that the two salaries must be equal.

The problem is that we have a **lost update.** The first transaction to commit, *T2,* overwrote Larry's salary as set by T1. **In** the serial order T2 followed by *T1,* Larry's salary should reflect T1's update rather than $T2$'s, but $T1$'s update is 'lost'.

16.3.4 Schedules Involving Aborted Transactions

We now extend our definition of serializability to include aborted trallsactions.[2] Intuitively, all actions of aborted transactions are to be undone, and we can therefore imagine that they were never carried out to begin with. Using this intuition, we extend the definition of a serializable schedule as follows: A **serializable schedule** over a set *S* of transactions is a schedule whose effect on any consistent database instance is guaranteed to be identical to that of some complete serial schedule over the set of *committed* transactions in S.

This definition of serializability relies on the actions of aborted transactions being undone completely, which may be impossible in some situations. For example, suppose that (1) an account transfer program T1 deducts $100 from account *A,* then (2) an interest deposit program T2 reads the current values of accounts *A* and *B* and adds 6% interest to each, then commits, and then (3) *T1* is aborted. The corresponding schedule is shown in Figure 16.5.

T1	*T2*
R(A)	
W(A)	
	R(A)
	W(A)
	R(B)
	W(B)
	Commit
Abort	

Figure 16.5 An Unrecoverable Schedule

[2] We must also consider incomplete transactions for a rigorous discussion of system failures, because transactions that are active when the system fails are neither aborted nor committed. However, system recovery usually begins by aborting all active transactions. and for our informal discussion, considering schedules involving committed and aborted transactions is sufficient.

Now, T2 has read a value for *A* that should never have been there. (Recall that aborted transactions' effects are not supposed to be visible to other transactions.) If *T2* had not yet committed, we could deal with the situation by *cascading* the abort of TI and also aborting *T2;* this process recursively aborts any transaction that read data written by *T2,* and so on. But *T2* has already committed, and so we cannot undo its actions. We say that such a schedule is *unrecoverable.* In a recoverable schedule, transactions commit only after (and if!) all transactions whose changes they read commit. If transactions read only the changes of committed transactions, not only is the schedule recoverable, but also aborting a transaction can be accomplished without cascading the abort to other transactions. Such a schedule is said to avoid cascading aborts.

There is another potential problem in undoing the actions of a transaction. Suppose that a transaction *T2* overwrites the value of an object *A* that has been modified by a transaction *T1,* while *T1* is still in progress, and *T1* subsequently aborts. All of *T1's* changes to database objects are undone by restoring the value of any object that it modified to the value of the object before *T1's* changes. (We look at the details of how a transaction abort is handled in Chapter 18.) When *T1* is aborted and its changes are undone in this manner, *T2's* changes are lost as well, even if *T2* decides to commit. So, for example, if *A* originally had the value 5, then was changed by *T1* to 6, and by *T2* to 7, if *T1* now aborts, the value of *A* becomes 5 again. Even if *T2* commits, its change to *A* is inadvertently lost. A concurrency control technique called Strict 2PL, introduced in Section 16.4, can prevent this problem (as discussed in Section 17.1).

16.4 LOCK-BASED CONCURRENCY CONTROL

A DBMS must be able to ensure that only serializable, recoverable schedules are allowed and that no actions of committed transactions are lost while undoing aborted transactions. A DBMS typically uses a *locking protocol* to achieve this. A lock is a small bookkeeping object associated with a database object. A locking protocol is a set of rules to be followed by each transaction (and enforced by the DBlVIS) to ensure that, even though actions of several transactions might be interleaved, the net effect is identical to executing all transactions in sOlne serial order. Different locking protocols use different types of locks, such as shared locks or exclusive locks, as we see next, when we discuss the Strict 2PL protocol.

16.4.1 Strict Two-Phase Locking (Strict 2PL)

The most widely used locking protocol, called *Strict Two-Phase Locking*, or *Strict 2PL*, has two rules. The first rule is

> 1. If a transaction T wants to *read* (respectively, *modify*) an object, it first requests a shared (respectively, exclusive) lock on the object.

Of course, a transaction that has an exclusive lock can also read the object; an additional shared lock is not required. A transaction that requests a lock is suspended until the DBMS is able to grant it the requested lock. The DBMS keeps track of the locks it has granted and ensures that if a transaction holds an exclusive lock on an object, no other transaction holds a shared or exclusive lock on the same object. The second rule in Strict 2PL is

> 2. All locks held by a transaction are released when the transaction is completed.

Requests to acquire and release locks can be automatically inserted into transactions by the DBMS; users need not worry about these details. eWe discuss how application programmers can select properties of transactions and control locking overhead in Section 16.6.3.)

In effect, the locking protocol allows only 'safe' interleavings of transactions. If two transactions access completely independent parts of the database, they concurrently obtain the locks they need and proceed merrily on their ways. On the other band, if two transactions access the same object, and one wants to modify it, their actions are effectively ordered serially·all actions of one of these transactions (the one that gets the lock on the common object first) are completed before (this lock is released and) the other transaction can proceed.

We denote the action of a transaction T requesting a shared (respectively, exclusive) lock on object 0 as $5T(0)$ (respectively, $XT(O))$ and omit the subscript denoting the transaction when it is clear from the context. As an example, consider the schedule shown in Figure 16.4. This interleaving could result in a state that cannot result from any serial execution of the three transactions. For instance, $T1$ could change A from 10 to 20, then $T2$ (which reads the value 20 for $A)$ could change B from 100 to 200, and then $T1$ would read the value 200 for B. If run serially, either T1 or $T2$ would execute first, and read the values 10 for A and 100 for B: Clearly, the interleaved execution is not equivalent to either serial execution.

If the Strict 2PL protocol is used, such interleaving is disallowed. Let us see why. Assuming that the transactions proceed at the same relative speed as

before, *T1* would obtain an exclusive lock on A first and then read and write A (Figure 16.6). Then, $T2$ would request a lock on A. However, this request

T1	*T2*
X(A)	
R(A)	
W(A)	

Figure 16.6 Schedule Illustrating Strict 2PL

cannot be granted until *T1* releases its exclusive lock on A, and the DBMS therefore suspends *T2*. *T1* now proceeds to obtain an exclusive lock on B, reads and writes B, then finally commits, at which time its locks are released. *T2's* lock request is now granted, and it proceeds. In this example the locking protocol results in a serial execution of the two transactions, shown in Figure 16.7.

T1	*T2*
X(A)	
R(A)	
W(A)	
X(B)	
R(B)	
W(B)	
Commit	
	X(A)
	R(A)
	W(A)
	X(B)
	R(B)
	W(B)
	Commit

Figure 16.7 Schedule Illustrating Strict 2PL with Serial Execution

In general, however, the actions of different transactions could be interleaved. As an example, consider the interleaving of two transactions shown in Figure 16.8, which is permitted by the Strict 2PL protocol.

It can be shown that the Strict 2PL algorithm allows only serializable schedules. None of the anomalies discussed in Section 16.3.3 can arise if the DBMS implements Strict 2PL.

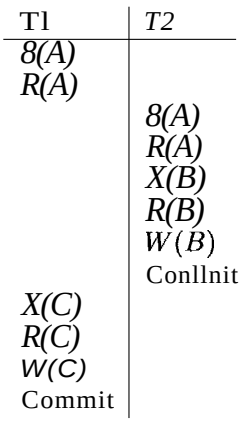


Figure 16.8 Schedule Following Strict 2PL with Interleaved Actions

16.4.2 Deadlocks

Consider the following example. Transaction *T1* sets an exclusive lock on object *A*, *T2* sets an exclusive lock on *B*, *T1* requests an exclusive lock on *B* and is queued, and *T2* requests an exclusive lock on *A* and is queued. Now, *T1* is waiting for *T2* to release its lock and *T2* is waiting for *T1* to release its lock. Such a cycle of transactions waiting for locks to be released is called a **deadlock.** Clearly, these two transactions will make no further progress. Worse, they hold locks that may be required by other transactions. The DBMS must either prevent or detect (and resolve) such deadlock situations; the common approach is to detect and resolve deadlocks.

A simple way to identify deadlocks is to use a timeout mechanism. If a transaction has been waiting too long for a lock, we can assume (pessimistically) that it is in a deadlock cycle and abort it. We discuss deadlocks in more detail in Section 17.2.

16.5 PERFORMANCE OF LOCKING

Lock-based schemes are designed to resolve conflicts between transactions and use two basic mechanisms: *blocking* and *aborting.* Both mechanisrns involve a performance penalty: Blocked transactions may hold locks that force other transactions to wait, and aborting and restarting a transaction obviously wastes the work done thus far by that transaction. A deadlock represents an extreme instance of blocking in which a set of transactions is forever blocked unless one of the deadlocked transactions is aborted by the DBMS.

In practice, fewer than 1% of transactions are involved in a deadlock, and there are relatively few aborts. Therefore, the overhead of locking comes primarily from delays due to blocking.[3] Consider how blocking delays affect throughput. The first few transactions are unlikely to conflict, and throughput rises in proportion to the number of active transactions. As more and more transactions execute concurrently on the same number of database objects, the likelihood of their blocking each other goes up. Thus, delays due to blocking increase with the number of active transactions, and throughput increases more slowly than the number of active transactions. In fact, there comes a point when adding another active transaction actually reduces throughput; the new transaction is blocked and effectively competes with (and blocks) existing transactions. We say that the system thrashes at this point, which is illustrated in Figure 16.9.

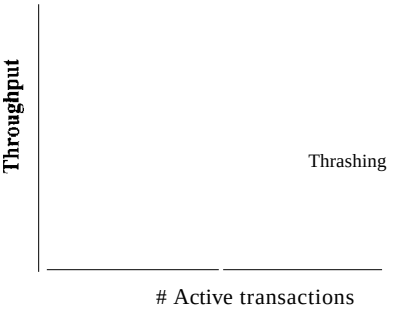


Figure 16.9 Lock Thrashing

If a database system begins to thrash, the database administrator should reduce the number of transactions allowed to run concurrently. Empirically, thrashing is seen to occur when 30% of active transactions are blocked, and a DBA should monitor the fraction of blocked transactions to see if the system is at risk of thrashing.

Throughput can be increased in three ways (other than buying a faster system):

- By locking the smallest sized objects possible (reducing the likelihood that two transactions need the same lock).
- By reducing the time that transaction hold locks (so that other transactions are blocked for a shorter time).

[3] Many common deadlocks can be avoided using a technique called *lock downgrade8,* implemented in most cOlnmercial systems (Section 17.3).

- By reducing **hot** spots. A hot spot is a databa.ge object that is frequently accessed and modified, and causes a lot of blocking delays. Hot spots can significantly affect performance.

The granularity of locking is largely determined by the database system's implementation of locking, and application programmers and the DBA have little control over it. We discuss how to improve performance by minimizing the duration locks are held and using techniques to deal with hot spots in Section 20.10.

16.6 TRANSACTION SUPPORT IN SQL

We have thus far studied transactions and transaction management using an abstract model of a transaction as a sequence of read, write, and abort/commit actions. We now consider what support SQL provides for users to specify transaction-level behavior.

16.6.1 Creating and Terminating Transactions

A transaction is automatically started when a user executes a statement that accesses either the database or the catalogs, such as a SELECT query, an UPDATE command, or a CREATE TABLE statement.[4]

Once a transaction is started, other statements can be executed as part of this transaction until the transaction is terminated by either a COMMIT command or a ROLLBACK (the SQL keyword for abort) command.

In SQL:1999, two new features are provided to support applications that involve long-running transactions, or that must run several transactions one after the other. To understand these extensions, recall that all the actions of a given transaction are executed in order, regardless of how the actions of different transactions are interleaved. We can think of each transaction as a sequence of steps.

The first feature, called a savepoint, allows us to identify a point in a transaction and selectively roll back operations carried out after this point. This is especially useful if the transaction carries out what-if kinds of operations, and wishes to undo or keep the changes based on the results. This can be accomplished by defining savepoints.

4Some SQL statements ..... e.g., the CONNECT statement, which connects an application program to a database server do not require the creation of a transaction.

SQL:1999 Nested Transactions: The concept of a transaction as an atomic sequence of actions has been extended in SQL:1999 thrQugh the introduction of the *savepoint* feature. This allows parts of a transaction to be selectively rolled back. The introduction of savepoints represents the first SQL support for the concept of nested transactions, which have been extensively studied in the research community. The idea is that a transaction can have several nested subtransactions, each of which can be selectively rolled back. Savepoints snpport a simple form of one-level nesting.

In a long-running transaction, we may want to define a series of savepoints. The savepoint command allows us to give each savepoint a name:

SAVEPDINT *(savepoint name)*

A subsequent rollback command can specify the savepoint to roll back to

ROLLBACK TO SAVEPDINT *(savepoint name)*

If we define three savepoints A, B, and C in that order, and then rollback to A, all operations since A are undone, including the creation of savepoints B and C. Indeed, the savepoint A is itself undone when we roll hack to it, and we must re-establish it (through another savepoint conunand) if we wish to be able to roll back to it again. From a locking standpoint, locks obtained after savepoint A can be released when we roll back to A.

It is instructive to compare the use of savepoints with the alternative of executing a series of transactions (i.e., treat all operations in between two consecutive savepoints as a new transaction). The savepoint mechanism offers two advantages. First, we can roll back over several savepoints. In the alternative approach, we can roll back only the most recent transaction, which is equivalent to rolling back to the most recent savepoint. Second, the overhead of initiating several transactions is avoided.

Even with the use of savepoints, certain applications might require us to run several transactions one after the other. To minimize the overhead in such situations, SQL:1999 introduces another feature, called chained transactions, \Ve can cornmit or roll back a transaction and immediately initiate another transaction. This is done by using the optional keywords AND CHAIN in the COMMIT and ROLLBACK statements.

16.6.2 What Should We Lock?

Until now, we have discussed transactions and concurrency control in tenus of an abstract model in which a database contains a fixed collection of objects, and each transaction is a series of read and write operations on individual objects. An important question to consider in the context of SQL is what the DBMS should treat as an *object* when setting locks for a given SQL statement (that is part of a transaction).

Consider the following query:

```
SELECT  S.rating, MIN  (S.age)
FROM    Sailors S
WHERE   S.rating = 8
```

Suppose that this query runs as part of transaction *T1* and an SQL statement that modifies the age of a given sailor, say Joe, with *rating=8* runs as part of transaction *T2*. What 'objects' should the DBMS lock when executing these transactions? Intuitively, we must detect a conflict between these transactions.

The DBMS could set a shared lock on the entire Sailors table for *T1* and set an exclusive lock on Sailors for *T2,* which would ensure that the two transactions are executed in a serializable manner. However, this approach yields low concurrency, and we can do better by locking smaller objects, reflecting what each transaction actually accesses. Thus, the DBMS could set a shared lock on every row with *rating=8* for transaction *T1* and set an exclusive lock on just the row for the modified tuple for transaction *T2*. Now, other read-only transactions that do not involve *rating=8* rows can proceed without waiting for *T1* or *T2*.

As this example illustrates, the DBMS can lock objects at different **granularities**: We can lock entire tables or set row-level locks. The latter approach is taken in current systems because it offers much better performance. **In** practice, while row-level locking is generally better, the choice of locking granularity is complicated. For example, a transaction that examines several rows and modifies those that satisfy some condition might be best served by setting shared locks on the entire table and setting exclusive locks on those rows it wants to lllodify. We diseuss this issue further in Section 17.5.3.

A second point to note is that SQL statements conceptually access a collection of rows described by a *selection predicate.* In the preceding example, transaction *T1* accesses all rows with *rating=8.* We suggested that this could be dealt with by setting shared locks on all rows in Sailors that had *rating=8.* Unfortunately, this is a little too silnplistic. To sec why, consider an SQL statelnent that inserts

a new sailor with *rating=8* and runs as transaction *T3*. (Observe that this example violates our assumption of a fixed number of objects in the database, but we must obviously deal with such situations in practice.)

Suppose that the DBJ\iIS sets shared locks on every existing Sailors row with *rating=8* for Tl. This does not prevent transaction *T3* from creating a brand new row with *rating=8* and setting an exclusive lock on this row. If this new row has a smaller *age* value than existing rows, *TI* returns an answer that depends on when it executed relative to *T2*. However, our locking scheme imposes no relative order on these two transactions.

This phenomenon is called the **phantom** problem: A transaction retrieves a collection of objects (in SQL terms, a collection of tuples) twice and sees different results, even though it does not modify any of these tuples itself. To prevent phantoms, the DBMS must conceptually lock *all possible* rows with *rating=8* on behalf of Tl. One way to do this is to lock the entire table, at the cost of low concurrency. It is possible to take advantage of indexes to do better, as we will see in Section 17.5.1, but in general preventing phantoms can have a significant impact on concurrency.

It may well be that the application invoking *T1* can accept the potential inaccuracy due to phantoms. If so, the approach of setting shared locks on existing tuples for *TI* is adequate, and offers better performance. SQL allows a programmer to make this choice---and other similar choices'--explicitly, as we see next.

16.6.3 Transaction Characteristics in SQL

In order to give programmers control over the locking overhead incurred by their transactions, SQL allows them to specify three characteristics of a transaction: access mode, diagnostics size, and isolation level. The **diagnostics** size determines the number of error conditions that can be recorded; we will not discuss this feature further.

If the access **mode** is READ ONLY, the transaction is not allowed to modify the database. Thus, INSERT, DELETE, UPDATE, and CREATE comlnands cannot be executed. If we have to execute one of these commands, the access mode should be set to READ WRITE. For transactions with READ ONLY access mode. only shared locks need to be obtained, thereby increasing concurrency.

The **isolation level** controls the extent to which a given transaction is exposed to the actions of other transactions executing concurrently. By choosing one of four possible isolation level settings, a user can obtain greater concur-

rencyat the cost of increasing the transaction's exposure to other transactions' uncommitted changes.

Isolation level choices are READ UNCOMMITTED, READ COMMITTED, REPEATABLE READ, and SERIALIZABLE. The effect of these levels is summarized in Figure 16.10. In this context, *dirty read* and *unrepeatable read* are defined as usuaL

Level	Dirty Read	Unrepeatable Read	
READ UNCOMMITTED	Maybe	Maybe	Maybe
READ COMMITTED	No	Maybe	Maybe
REPEATABLE READ	No	No	Maybe
SERIALIZABLE	No	No	No

Figure 16.10 Transaction Isolation Levels in SQL-92

The highest degree of isolation from the effects of other transactions is achieved by setting the isolation level for a transaction T to SERIALIZABLE. This isolation level ensures that T reads only the changes made by committed transactions, that no value read or written by T is changed by any other transaction until T is complete, and that if T reads a set of values based on some search condition, this set is not changed by other transactions until T is complete (i.e., T avoids the phantom phenomenon).

In terms of a lock-based implementation, a SERIALIZABLE transaction obtains locks before reading or writing objects, including locks on sets of objects that it requires to be unchanged (see Section 17.5.1) and holds them until the end, according to Strict 2PL.

REPEATABLE READ ensures that T reads only the changes made by committed transactions and no value read or written by T is changed by any other transaction until T is complete. However, T could experience the phantom phenomenon; for example, while T examines all Sailors records with $rating=1$, another transaction might add a new such Sailors record, which is missed by T.

A REPEATABLE READ transaction sets the same locks as a SERIALIZABLE transaction, except that it does not do index locking; that is, it locks only individual objects, not sets of objects. We discuss index locking in detail in Section 17.5.1.

READ COMMITTED ensures that T reads only the changes made by committed transactions, and that no value written by T is changed by any other transaction until T is complete. However, a value read by T may well be modified by

another transaction while T is still in progress, and T is exposed to the phantom problem.

A READ COMMITTED transaction obtains exclusive locks before writing objects and holds these locks until the end. It also obtains shared locks before reading objects, but these locks are released immediately; their only effect is to guarantee that the transaction that last modified the object is complete. (This guarantee relies on the fact that *every* SQL transaction obtains exclusive locks before writing objects and holds exclusive locks until the end.)

A READ UNCOMMITTED transaction T can read changes made to an object by an ongoing transaction; obviously, the object can be changed further while T is in progress, and T is also vulnerable to the phantom problem.

A READ UNCOMMITTED transaction does not obtain shared locks before reading objects. This mode represents the greatest exposure to uncommitted changes of other transactions; so much so that SQL prohibits such a transaction from making any changes itself-a READ UNCOMMITTED transaction is required to have an access mode of READ ONLY. Since such a transaction obtains no locks for reading objects and it is not allowed to write objects (and therefore never requests exclusive locks), it never makes any lock requests.

The SERIALIZABLE isolation level is generally the safest and is recommended for most transactions. Some transactions, however, can run with a lower isolation level, and the smaller number of locks requested can contribute to improved system performance. For example, a statistical query that finds the average sailor age can be run at the READ COMMITTED level or even the READ UNCOMMITTED level, because a few incorrect or missing values do not significantly affect the result if the number of sailors is large.

The isolation level and access mode can be set using the SET TRANSACTION command. For example, the following command declares the current transaction to be SERIALIZABLE and READ ONLY:

```
SET TRANSACTION ISOLATION LEVEL SERIALIZABLE READ ONLY
```

When a transaction is started, the default is SERIALIZABLE and READ WRITE.

16.7 INTRODUCTION TO CRASH RECOVERY

The recovery manager of a DBMS is responsible for ensuring transaction *atomicity* and *durability*. It ensures atomicity by undoing the actions of transactions that do not commit, and durability by making sure that all actions of

committed transactions survive **systenl crashes,** (e.g., a core dump caused by a bus error) and **media failures** (e.g., a disk is corrupted).

\\Then a DBMS is restarted after crashes. the recovery manager is given control and must bring the database to a consistent state. The recovery manager is also responsible for undoing the actions of an aborted transaction. To see what it takes to implement a recovery manager, it is necessary to understand what happens during normal execution.

The **transaction manager** of a DBMS controls the execution of transactions. Before reading and writing objects during normal execution, locks must be acquired (and released at some later time) according to a chosen locking protocol.[5] For simplicity of exposition, we make the following assumption:

Atomic Writes: Writing a page to disk is an atomic action.

This implies that the system does not crash while a write is in progress and is unrealistic. **In** practice, disk writes do not have this property, and steps must be taken during restart after a crash (Section 18.6) to verify that the most recent write to a given page was completed successfully, and to deal with the consequences if not.

16.7.1 Stealing Frames and Forcing Pages

\Vith respect to writing objects, two additional questions arise:

1. Can the changes made to an object 0 in the buffer pool by a transaction *T* be written to disk before *T* commits? Such writes are executed when another transaction wants to bring in a page and the buffer manager chooses to replace the frame containing 0; of course, this page must have been unpinned by *T*. **If** such writes are allowed, we say that a **steal** approach is used. (Informally, the second transaction 'steals' a frame from *T.)*
2. When a transaction cOl1units, must we ensure that all the changes it has made to objects in the buffer pool are immediately forced to disk? **If** so. we say that a **force** approach is used.

From the standpoint of implementing a recovery manager, it is simplest to use a buffer manager with a no-steaL force approach. **If** a no-steal approach is used, we do not have to undo the changes of an aborted transaction (because these changes have not been written to disk), and if a force approach is used, we do

[5] A concurrency control technique that does not involve locking could be used instead, but we assume that locking is llsed.

not have to redo the changes of a committed transaction if there is a subsequent crash (because all these changes are guaranteed to have been written to disk at commit time).

However, these policies have important drawbacks. The no-steal approach assumes that all pages modified by ongoing transactions can be accommodated in the buffer pool, and in the presence of large transactions (typically run in batch mode, e.g., payroll processing), this assumption is unrealistic. The force approach results in excessive page I/O costs. If a highly used page is updated in succession by 20 transactions, it would be written to disk 20 times. With a no-force approach, on the other hand, the in-memory copy of the page would be successively modified and written to disk just once, reflecting the effects of all 20 updates, when the page is eventually replaced in the buffer pool (in accordance with the buffer manager's page replacement policy).

For these reasons, most systems use a steal, no-force approach. Thus, if a frame is dirty and chosen for replacement, the page it contains is written to disk even if the modifying transaction is still active *(steal);* in addition, pages in the buffer pool that are modified by a transaction are not forced to disk when the transaction commits *(no-force).*

16.7.2 Recovery-Related Steps during Normal Execution

The recovery manager of a DBMS maintains some information during normal execution of transactions to enable it to perform its task in the event of a failure. In particular, a log of all modifications to the database is saved on stable storage, which is guaranteed[6] to survive crashes and media failures. Stable storage is implemented by maintaining multiple copies of information (perhaps in different locations) on nonvolatile storage devices such as disks or tapes.

As discussed earlier in Section 16.7, it is important to ensure that the log entries describing a change to the database are written to stable storage *before* the change is made; otherwise, the system might crash just after the change, leaving us without a record of the change. (Recall that this is the Write-Ahead Log, or WAL, property.)

The log enables the recovery manager to undo the actions of aborted and incomplete transactions and redo the actions of committed transactions. For example, a transaction that committed before the crash may have made updates

(jNothing in life is really guaranteed except death and taxes. However, we can reduce the chance of log failure to be vanishingly small by taking steps such as duplexing the log and storing the copies in different secure locations.

Tuning the Recovery Subsystem: DBMS performance can be greatly affected by the overhead imposed by the recoverysubsystem. A DBA can take several steps to tune this subsystem, such as correctlysizing the log and how it is managed on disk, controlling the rate at which buffer pages are forced to disk, choosing a good frequency for checkpointing, and so forth.

to a copy (of a database object) in the buffer pool, and this change may not have been written to disk before the crash, because of a no-force approach. Such changes must be identified using the log and written to disk. Further, changes of transactions that did not commit prior to the crash might have been written to disk because of a steal approach. Such changes must be identified using the log and then undone.

The amount of work involved during recovery is proportional to the changes made by committed transactions that have not been written to disk at the time of the crash. To reduce the time to recover from a crash, the DBMS periodically forces buffer pages to disk during normal execution using a background process (while making sure that any log entries that describe changes these pages are written to disk first, i.e., following the WAL protocol). A process called *checkpointing,* which saves information about active transactions and dirty buffer pool pages, also helps reduce the time taken to recover from a crash. Checkpoints are discussed in Section 18.5.

16.7.3 Overview of ARIES

ARIES is a recovery algorithm that is designed to work with a steal, no-force approach. When the recovery manager is invoked after a crash, restart proceeds in three phases. In the Analysis phase, it identifies dirty pages in the buffer pool (i.e., changes that have not been written to disk) and active transactions at the time of the crash. In the Redo phase, it repeats all actions, starting from an appropriate point in the log, and restores the database state to what it was at the time of the crash. Finally, in the Undo phase, it undoes the actions of transactions that did not commit, so that the database reflects only the actions of committed transactions. The ARIES algorithm is discussed further in Chapter 18.,

16.7.4 Atomicity: Implementing Rollback

It is important to recognize that the recovery subsystem is also responsible for executing the ROLLBACK command, which aborts a single transaction. Indeed,

the logic (and code) involved in undoing a single transaction is identical to that used during the Undo phase in recovering from a system crash. All log records for a given transaction are organized in a linked list and can be efficiently accessed in reverse order to facilitate transaction rollback.

16.8 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are the ACID properties? Define *atomicity, consistency, isolation,* and *durability* and illustrate them through examples, **(Section 16.1)**
- Define the terms *transaction, schedule, complete schedule,* and *seTial schedule.* **(Section 16.2)**
- Why does a DBMS interleave concurrent transactions? **(Section 16.3)**
- When do two actions on the same data object *conflict?* Define the anomalies that can be caused by conflicting actions *(dirty reads, unrepeatable reads, lost updates).* **(Section 16.3)**
- What is a *serializable schedule?* What is a *Tecoverable schedule?* What is a schedule that *avoids cascading abor'ts?* What is a *strict schedule?* **(Section 16.3)**
- What is a *locking protocol?* Describe the *Strict Two-Phase Locking (StTict 2PL)* protocol. What can you say about the schedules allowed by this protocol? **(Section 16.4)**
- What overheads are associated with lock-based concurrency control? Discuss *blocking* and *aborting* overheads specifically and explain which is more important in practice. **(Section 16.5)**
- What is thrashing? What should a DBA do if the system thrashes? **(Section 16.5)**
- How can throughput be increased? **(Section 16.5)**
- How are transactions created and terminated in SQL? What are savepoints? What are chained transactions? Explain why savepoints and chained transactions are useful. **(Section 16.6)**
- What are the considerations in determining the locking granularity when executing SQL statements? What is the phantom problem? What irnpact does it have on performance? **(Section 16.6.2)**

- What transaction characteristics can a programmer control in SQL? Discuss the different *access modes* and *isolat'ion levels* in particular. What issues should be considered in selecting an access mode and an isolation level for a transaction? (Section 16.6.3)
- Describe how different isolation levels are implemented in terms of the locks that are set. 'What can you say about the corresponding locking overheads? (Section 16.6.3)
- What functionality does the *recovery manager* of a DBMS provide? What does the *transaction manager* do? (Section 16.7)
- Describe the *steal* and *force* policies in the context of a buffer manager. What policies are used in practice and how does this affect recovery? (Section 16.7.1)
- What recovery-related steps are taken during normal execution? What can a DBA control to reduce the time to recover from a crash? (Section 16.7.2)
- How is the log used in transaction rollback and crash recovery? (Sections 16.7.2, 16.7.3, and 16.7.4)

EXERCISES

Exercise 16.1 Give brief answers to the following questions:

1. What is a transaction? In what ways is it different from an ordinary program (in a language such as C)?
2. Define these terms: *atomicity, consistency, isolation, durability, schedule, blind write, dirty read, unrepeatable read, serializable schedule, recoverable schedule, avoidsvcascading-aborts schedule.*
3. Describe Strict 2PL.
4. What is the phantom problem? Can it occur in a database where the set of database objects is fixed and only the values of objects can be changed?

Exercise 16.2 Consider the following actions taken by transaction 1'1 on database objects *X* and *Y:*

R(X), W(X),R(Y), W(Y)

1. Give an example of another transaction 1'2 that, if run concurrently to transaction *T* without some form of concurrency control, could interfere with 1'1.
2. Explain how the use of Strict 2PL would prevent interference between the two transactions.
3. Strict 2PL is lIsed in many database systems. Give two reasons for its popularity.

Exercise 16.3 Consider a database with objects X and Y and assume that there are two transactions $T1$ and $T2$. Transaction $T1$ reads objects X and Y and then writes object X. Transaction $T2$ reads objects X and Y and then writes objects X and Y.

1. Give an example schedule with actions of transactions $T1$ and $T2$ on objects X and Y that results in a write-read conflict.
2. Give an example schedule with actions of transactions $T1$ and $T2$ on objects X and Y that results in a read-write conflict.
3. Give an example schedule with actions of transactions $T1$ and $T2$ on objects X and Y that results in a write-write conflict.
4. For each of the three schedules, show that Strict 2PL disallows the schedule.

Exercise 16.4 We call a transaction that only reads database object a read-only transaction, otherwise the transaction is called a read-write transaction. Give brief answers to the following questions:

1. What is lock thrashing and when does it occur?
2. What happens to the database system throughput if the number of read-write transactions is increased?
3. What happens to the datbase system throughput if the number of read-only transactions is increased?
4. Describe three ways of tuning your system to increase transaction throughput.

Exercise 16.5 Suppose that a DBMS recognizes *increment*, which increments an integer-valued object by 1, and *decrement* as actions, in addition to reads and writes. A transaction that increments an object need not know the value of the object; increment and decrement are versions of blind writes. In addition to shared and exclusive locks, two special locks are supported: An object must be locked in I mode before incrementing it and locked in D mode before decrementing it. An I lock is compatible with another I or D lock on the same object, but not with S and X locks.

1. Illustrate how the use of I and D locks can increase concurrency. (Show a schedule allowed by Strict 2PL that only uses S and X locks. Explain how the use of I and D locks can allow more actions to be interleaved, while continuing to follow Strict 2PL.)
2. Informally explain how Strict 2PL guarantees serializability even in the presence of I and D locks. (Identify which pairs of actions conflict, in the sense that their relative order can affect the result, and show that the use of S, X, I, and D locks according to Strict 2PL orders all conflicting pairs of actions to be the same as the order in some serial schedule.)

Exercise 16.6 Answer the following questions: SQL supports four isolation-levels and two access-modes, for a total of eight combinations of isolation-level and access-mode. Each combination impiicitly defines a class of transactions; the following questions refer to these eight classes:

1. Consider the four SQL isolation levels. Describe which of the plHmomena can occur at each of these isolation levels: *dirty read, unrepeatable read, phantom problem.*
2. For each of the four isolation levels, give examples of transactions that could be run safely at that level.
3. Why does the access mode of a transaction matter?

Exercise 16.7 Consider the university enrollment database schema:

Student(snurn: integer, *snarne:* string, *majoT:* string, *level:* string, *age:* integer)
Class(*name:* string, *meets_at:* time, *room:* string, *fid:* integer)
Enrolled(snum: integer, *cname:* string)
Faculty(*fid:* integer, *fname:* string, *deptid:* integer)

The meaning of these relations is straightforward; for example, Enrolled has one record per student-class pair such that the student is enrolled in the class.

For each of the following transactions, state the SQL isolation level you would use and explain why you chose it.

1. Enroll a student identified by her *snum* into the class named 'Introduction to Database Systems'.
2. Change enrollment for a student identified by her *snum* from one class to another class,
3. Assign a new faculty member identified by his *fid* to the class with the least number of students.
4. For each class, show the number of students enrolled in the class.

Exercise 16.8 Consider the following schema:

Suppliers(sid: integer, *sname:* string, *addTess:* string)
Parts(pid: integer, *pname:* string, *coloT:* string)
Catalog(sid: integer, *pid:* integer, *cost:* real)

The Catalog relation lists the prices charged for parts by Suppliers.

For each of the following transactions, state the SQL isolation level that you would use and explain why you chose it.

1. A transaction that adds a new part to a supplier's catalog.
2. A transaction that increases the price that a supplier charges for a part.
3. A transaction that determines the total number of items for a given supplier.
4. A transaction that shows, for each part, the supplier that supplies the part at the lowest price.

Exercise 16.9 Consider a database with the following schema:

Suppliers(*sid:* integer, *sname:* string, *addTess:* string)
Parts(pid: integer, *pname:* string, *coloT:* string)
Catalog(*sid:* integer, *pid:* integer, *cost:* real)

The Catalog relation lists the prices charged for parts by Suppliers.

Consider three transactions *1'1,1'2,* and *1'3; 1'1* always has SQL isolation level SERIALIZABLE. We first run *1'1* concurrently with *1'2* and then we run *1'1* concurrently with *1'2* but we change the isolation level of *1'2* as specified below. Give a database instance and SQL statements for *1'1* and *1'2* such that result of running *1'2* with the first SQL isolation level is different from running 1'2 with the second SQL isolation level. Also specify the common schedule of *1'1* and *1'2* and explain why the results are different.

1. SERIALIZABLE versus REPEATABLE READ.
2. REPEATABLE READ versus READ COMMITTED.
3. READ COMMITTED versus READ UNCOMMITTED.

BIBLIOGRAPHIC NOTES

The transaction concept and some of its limitations are discussed in [332J. A formal transaction model that generalizes several earlier transaction models is proposed in [182].

Two-phase locking was introduced in [252], a fundamental paper that also discusses the concepts of transactions, phantoms, and predicate locks. Formal treatments of serializability appear in [92, 581].

Excellent in-depth presentations of transaction processing can be found in [90] and [770]. [338] is a classic, encyclopedic treatment of the subject.

17

CONCURRENCY CONTROL

- How does Strict 2PL ensure serializability and recoverability?
- How are locks implemented in a DBMS?
- What are lock conversions and why are they important?
- How does a DBMS resolve deadlocks?
- How do current systerns deal with the phantom problerrl?
- Why are specialized locking techniques used on tree indexes?
- How does multiple-granularity locking work?
- What is Optimistic concurrency control?
- What is Timestamp-Based concurrency control?
- What is Multiversion concurrency control?
- Key concepts: Two-phase locking (2PL), serializability, recoverability, precedence graph, strict schedule, view equivalence, view serializable, lock nlanager, lock table, transaction table, latch, convoy, lock upgrade, deadlock, waits-for graph, conservative 2PL, index locking, predicate locking, multiple-granularity locking, lock escalation, SQL isolation level, phantom problerrl, optirnistic concurrency control, Thornas Write Rule, recoverability

Pooh was sitting in his house one day, counting his pots of honey, when there carne a knock on the door.
"Fourteen," said Pooh. "Come in. Fourteen. Or was it fifteen? Bother. T'hat's rnuddled rnc."

"Hallo, Pooh," said Rabbit. "Halla, R,abbit. Fourteen, wasn't it?"
"What was?" "My pots of honey what I was counting."
"Fourteen, that's right."
"Are you sure?"
"No," said Rabbit. "Does it matter?"

—A.A. Milne, *The House at Pooh Comer*

In this chapter, we look at concurrency control in more detail. We begin by looking at locking protocols and how they guarantee various irnportant properties of schedules in Section 17.1. Section 17.2 is an introduction to how locking protocols are implemented in a DBMS. Section 17.3 discusses the issue of lock conversions, and Section 17.4 covers deadlock handling. Section 17.5 discusses three specialized locking protocols---for locking sets of objects identified by some predicate, for locking nodes in tree-structured indexes, and for locking collections of related objects. Section 17.6 examines some alternatives to the locking approach.

17.1 2PL, SERIALIZABILITY, AND RECOVERABILITY

In this section, we consider how locking protocols guarantee some important properties of schedules; namely, serializability and recoverability. Two schedules are said to be conflict equivalent if they involve the (sarne set of) actions of the same transactions and they order every pair of conflicting actions of two committed transactions in the sanle way.

As we saw in Section 16.3.3, two actions conflict if they operate on the same data object and at least one of them is a write. The outcome of a schedule depends only on the order of conflicting operations; we can interchange any pair of nonconflicting operations without altering the effect of the schedule on the database. If two schedules are conflict equivalent, it is easy to see that they have the same effect on a database. Indeed, because they order all pairs of conflicting operations in the same way, we can obtain one of thern frorn the other by repeatedly swapping pairs of nonconflicting actions, that is, by swapping pairs of actions whose relative order does not alter the outcome.

A schedule is conflict serializable if it is conflict equivalent to some serial schedule. Every conflict serializable schedule is serializable, if we assurne that the set of items in the database does not grow or shrink; that is, values can be nlodified but items are not added or deleted. We lllake this assurnption for now and consider its consequences in Section 17.5.1. However, sonle serializable schedules are not conflict serializable, as illustrated in Figure 17.1. This schedule is equivalent to executing the transactions serially in the order $T1$, $T2$,

T1	T2	T3
R(A)		
	W(A)	
	COllirnit	
W(A)		
COllirnit		
		W(A)
		Commit

Figure 1.7.1 Serializable Schedule That Is Not Conflict Serializable

T3, but it is not conflict equivalent to this serial schedule because the writes of *T1* and *T2* are ordered differently.

It is useful to capture all potential conflicts between the transactions in a schedule in a precedence graph, also called a serializability graph. The precedence graph for a schedule *S* contains:

- A node for each comnlitted transaction in *S.*
- An arc franl *Ti* to *Tj* if an action of *Ti* precedes and conflicts with one of *Tj*'s actions.

The precedence graphs for the schedules shown in Figures 16.7, 16.8, and 17.1 are shown in Figure 17.2 (parts a, b, and c, respectively).

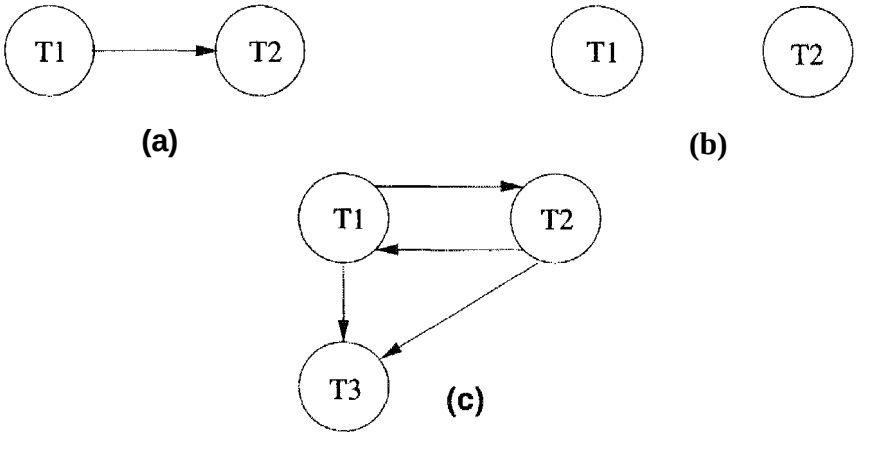


Figure 17.2 Examples of Precedence Graphs

'I'he Strict 2PL protocol (introduced in Section 16.4) allows only conflict serializable schedules, as is seen frcHu the following two results:

1. A schedule S' is conflict serializable if and only if its precedence graph is acyclic. (An equivalent serial schedule in this case is given by any topological sort over the precedence graph.)

2. Strict 2PL ensures t.hat the precedence graph for any schedule that it allows is acyclic.

A widely studied variant of Strict 2PL, called Two-Phase Locking (2PL), relaxes the second rule of Strict 2PL to allow transactions to release locks before the end, that is, before the comnlit or abort action. For 2PL, the second rule is replaced by the following rule:

> (2PL) (2) A transaction cannot request additional locks once it releases *any* lock.

Thus, every transaction has a 'growing' phase in which it acquires locks, followed by a 'shrinking' phase in which it releases locks.

It can be shown that even nonstrict 2PL ensures acyclicity of the precedence graph and therefore allows only conflict serializable schedules. Intuitively, an equivalent serial order of transactions is given by the order in which transactions enter their shrinking phase: If *T2* reads or writes an object written by *Tl*, *Tl* IllUSt have released its lock on the object before *T*'2 requested a lock on this object. Thus, *Tl* precedes *T2*. (A sirnilar argulnent shows that *Tl* precedes *T2* if 7'2 writes an object previously read by *Tl*. A forIllal proof of the claim would have to show that there is no cycle of transactions that 'precede' each other by this argurnent.)

A schedule is said to be strict if a value written by a transaction *T* is not read or overwritten by other transactions until *T* either aborts or eOlnrnits. Strict schedules are recoverable, do not require cascading aborts, and actions of aborted transactions can be undone by restoring the original values of lnodified objects. (See the last exaInple in Section 16.3.4.) Strict 2PL irnproves on 2PL by guaranteeing that every allowed schedule is strict in addition to being conflict serializable. The reason is that when a transaction *T* writes an object under Strict 2PL, it holds the (exclusive) lock until it conunits or aborts. Thus, no other transaction can see or rnodify this object until *T* is cornplete.

The reader is invited to revisit the exarnples in Section 16.3.3 to see how the corresponding schedules are disallowed by Strict 2PL and 2PL. Sirnilarly, it would be instructive to \vork out how the schedules for the exarnples in Section 16.3.4 are disallowed by Strict 2PL but not by 2PL.

17.1.1 View Serializability

Conflict serializability is sufficient but not necessary for serializability. A 1n01'e general sufficient condition is view serializability. Two schedules 81 and 82 over the saIne set of transactions—any transaction that appears in either 81 or 82 rnust also appear in the other----------are view equivalent under these conditions:

1. If *fTi* reads the initial value of object *A* in 81, it Blust also read the initial value of *A* in 82.
2. If *Ti* reads a value of *A* written by *Tj* in 81, it IllUst also read the value of *A* written by *Tj* in 82.
3. For each data object *A*, the transaction (if any) that perforlns the final write on *A* in 81 must also perform the final write on *A* in 82.

A schedule is view serializable if it is view equivalent to SaIne serial schedule. Every conflict serializable schedule is view serializable, although the converse is not true. For example, the schedule shown in Figure 17.1 is view serializable, although it is not conflict serializable. Incidentally, note that this exalnple contains blind writes. This is not a coincidence; it can be shown that any view serializable schedule that is not conflict serializable contains a blind write.

As we saw in Section 17.1, efficient locking protocols allow us to ensure that only conflict serializable schedules are allowed. Enforcing or testing vie\v serializability turns out to be lIluch lnore expensive, and the concept therefore has little practical use, although it increases our understanding of serializability.

17.2 INTRODUCTION TO LOCK MANAGEMENT

The part of the DBMS that keeps track of the locks issued to transactions is called the lock manager. The lock lnanager rnaintains a lock table, which is a hash table with the data object identifier as the key. The DBMS also Inaintains a descriptive entry for each transaction in a transaction table, and alIlong other things, the entry contains a pointer to a list of locks held by the transaction. This list is checked before requesting a lock, to ensure that a transaction does not request the saIne lock twice.

A lock table entry for an object—which can be a page, a record, and so on, depending on the DBMS---contains the following inforrnation: the nurnber of transactions currently holding a lock on the object (this can be rnore than one if the object is locked in shared rnode), the nature of the lock (shared or exclusive), and a pointer to a queue of lock requests.

17.2.1 Implementing Lock and Unlock Requests

According to the Strict 2PL protocol, before a transaction *T* reads or writes a database object 0, it must obtain a shared or exclusive lock on 0 and Inust hold on to the lock until it commits or aborts. When a transaction needs a lock on an object, it issues a lock request to the lock manager:

1. If a shared lock is requested, the queue of requests is ernpty, and the object is not currently locked in exclusive mode, the lock manager grants the lock and updates the lock table entry for the object (indicating that the object is locked in shared mode, and incrernenting the number of transactions holding a lock by one).
2. If an exclusive lock is requested and no transaction currently holds a lock on the object (which also implies the queue of requests is empty), the lock rnanager grants the lock and updates the lock table entry.
3. Otherwise, the requested lock cannot be immediately granted, and the lock request is added to the queue of lock requests for this object. The transaction requesting the lock is suspended.

When a transaction aborts or comrnits, it releases all its locks. When a lock on an object is released, the lock manager updates the lock table entry for the object and exarnines the lock request at the head of the queue for this object. If this request can now be granted, the transaction that made the request is woken up and given the lock. Indeed, if several requests for a shared lock on the object are at the front of the queue, all of these requests can now be granted together.

Note that if *TI* has a shared lock on 0 and 1'2 requests an exclusive lock, *T2's* request is queued. Now, if *T3* requests a shared lock, its request enters the queue behind that of *T2*, even though the requested lock is cornpatible with the lock held by TI. This rule ensures that *T2* does not *starve*, that is, wait indefinitely while a stream of other transactions acquire shared locks and thereby prevent *T2* frorn getting the exclusive lock for which it is waiting.

Atomicity of Locking and Unlocking

The irnplernentation of *lock* and *unlock* cornrnands rnust ensure that these are atomic operations. To ensure atornicity of these operations when several instances of the lock rnanager code can exccute concurrently, access to the lock table has to be guarded by an operating systern synchronization rnechanisrn such as a sernaphore.

To understand why, suppose that a transaction requests an exclusive lock. The lock manager checks and finds that no other transaction holds a lock on the object and therefore decides to grant the request. But, in the 11leantirne, another transaction rnight have requested and *received* a conflicting lock. To prevent this, the entire sequence of actions .in a lock request call (checking to see if the request can be granted, updating the lock table, etc.) must be irnplernented as an atornic operation.

Other Issues: Latches, Convoys

In addition to locks, which are held over a long duration, a DBMS also supports short-duration latches. Setting a latch before reading or writing a page ensures that the physical read or write operation is atomic; otherwise, two read/write operations rnight conflict if the objects being locked do not correspond to disk pages (the units of I/O). Latches are unset immediately after the physical read or write operation is cOlnpleted.

We concentrated thus far on how the DBMS schedules transactions based on their requests for locks. This interleaving interacts with the operating system's scheduling of processes' access to the CPU and can lead to a situation called a convoy, where most of the CPU cycles are spent on process switching. The problem is that a transaction *T* holding a heavily used lock may be suspended by the operating system. UntH *T* is resurned, every other transaction that needs this lock is queued. Such queues, called convoys, can quickly become very long; a convoy, once forrned, tends to be stable. Convoys are one of the drawbacks of building a DBMS on top of a general-purpose operating system with preeruptive scheduling.

17.3 LOCK CONVERSIONS

A transaction rnay need to acquire an exclusive lock on an object for which it already holds a shared lock. For exarnple, a SQL update statenlent could result in shared locks being set on each row in a table. If a row satisfies the condition (in the WHERE clause) for being updated, an exclusive lock must be obtained for that row.

Such a lock upgrade request lnust be handled specially by granting the exclusive lock illunediately if no other transaction holds a shared lock on the object and inserting the request at the front of the queue other\vise. The rationale for favoring the transaction thus is that it already 1101ds a shared lock on the object and queuing it behind. another transaction that wants an exclusive lock on the same object causes both a deadlock. UnfortunatelY,while favoring lock upgrades helps, it does not prevent deadlocks caused by two conflicting upgrade

requests. For exalnplc, if two transactions that hold a shared lock on an object both request an upgrade to an exclusive lock, this leads to a deadlock.

A better approach is to avoid the need for lock upgrades altogether by obtaining exclusive locks initially, and downgrading to a shared lock once it is clear that this is sufficient. In our exalnple of an SQL update statelnent, rows in a table are locked in exclusive rnode first. If a row does *not* satisfy the condition for being updated, the lock on the row is dnwngraded to a shared lock. Does the downgrade approach violate the 2PL requirernent? On the surface, it does, because downgrading reduces the locking privileges held by a transaction, and the transaction Illay go on to acquire other locks. However, this is a special case, because the transaction did nothing but read the object that it downgraded, even though it conservatively obtained an exclusive lock. We can safely expand our definition of 2PL from Section 17.1 to allow lock downgrades in the growing phase, provided that the transaction has not lnodified the object.

The downgrade approach reduces concurrency by obtaining write locks in some cases where they are not required. On the whole, however, it irnproves throughput by reducing deadlocks. This approach is therefore widely used in current commercial systems. Concurrency can be increased by introducing a new kind of lock, called an update lock, that is cornpatible with shared locks but not other update and exclusive locks. By setting an update lock initially, rather than exclusive locks, we prevent conflicts with other read operations. Once we are sure we need not update the object, we can downgrade to a shared lock. If we need to update the object, we rnust first upgrade to an exclusive lock. This upgrade does not lead to a deadlock because no other transaction can have an upgrade or exclusive lock on the object.

17.4 DEALING WITH DEADLOCKS

Deadlocks tend to be rare and typically involve very few transactions. In practice, therefore, database systerns periodically check for deadlocks. When a transaction Ti is suspended because a lock that it requests cannot be granted, it rnust wait until all transactions Tj that currently hold conflicting locks release thcrn. The lock rnanager rnaintains a structure called a waits-for graph to detect deadlock cycles. The nodes correspond to active transactions, and there is an arc frolnTi to Tj if (and only if)Ti is \vaiting for Tj to release a lock. The lock rnanagcr adds edges to this graph when it queues lock requests and rernoves edges when it gra,nts lock requests.

Consider the schedule shown in F'igure 17.:3, The last step, shown below the line, creates a cycle in the \vaits-for graph. Figure 17.4 shows the ·waits-for graph before and after this step.

TI	*T2*	*T3*	*T4*
S(A)			
$R(A)$			
	X(B)		
	$W(B)$		
8(B)			
		8(C)	
		R(C)	
	X(C)		
			X(B)
		X(A)	

Figure 17.3 Schedule Illustrating Deadlock

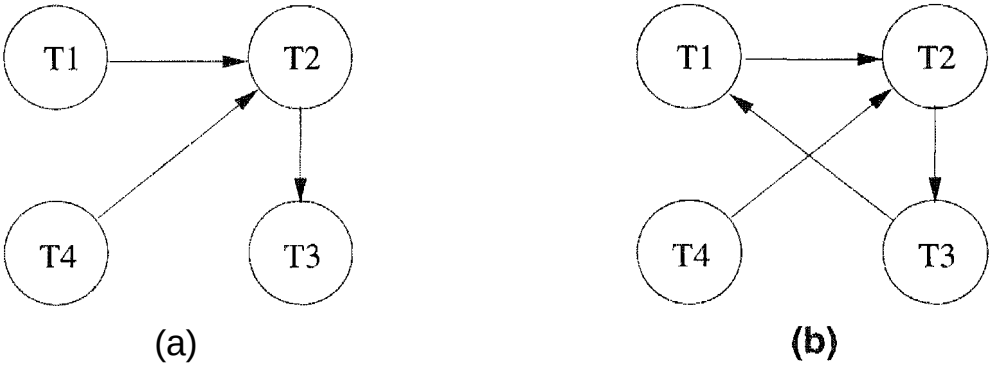


Figure 17.4 Waits-for Graph Before and After Deadlock

Observe that the \vaits-for graph describes all active transactions, some of which eventually abort. If there is an edge froIn *Ti* to *Tj* in the waits-for graph, and both *Ti* and *Tj* eventually commit, there is an edge in the opposite direction (froIn *I'j* to *Ti)* in the precedence graph (which involves only cOlluuitted transactions).

The waits-for graph is periodically checked for cycles, which indicate deadlock. A deadlock is resolved by aborting a transaction that is on a cycle and releasing its locks; this action allows SOlne of the waiting transactions to proceed. The choice of which transaction to abort can be made using several criteria: the one with the fewest locks, the one that has done the least work, the one that is farthest from completion, and so all. Further, a transaction might have been repeatedly restarted; if so, it should eventually be favored during deadlock detection and allowed to complete.

A simple alternative to maintaining a waits-for graph is to identify deadlocks through a timeout mechanism: If a transaction has been waiting too long for a lock, we assume (pessiruistically) that it is in a deadlock cycle and abort it.

17.4.1 Deadlock Prevention

Elnpirical results indicate that deadlocks are relatively infrequent, and detection-based schemes work well in practice. However, if there is a high level of contention for locks and therefore an increased likelihood of deadlocks, prevention-based schelnes could perform better. We can prevent deadlocks by giving each transaction a priority and ensuring that lower-priority transactions are not allowed to wait for higher-priority transactions (or vice versa). One way to assign priorities is to give each transaction a timestamp when it starts up. The lower the timestamp, the higher is the transaction's priority; that is, the oldest transaction has the highest priority.

If a transaction *Ti* requests a lock and transaction *Tj* holds a conflicting lock, the lock lnanager can use one of the following two policies:

- Wait-die: If *Ti* has higher priority, it is allowed to wait; otherwise, it is aborted.
- Wound-wait: If *Ti* has higher priority, abort *Tj*; otherwise, *Ti* waits.

In the \vait-die scherne, lower-priority transactions can never wait for higher-priority transactions. In the wound-wait scherne, higher-priority transactions never wait for lower-priority transactions. In either ease, no deadlock cyc.le develops.

A subtle point is that we nlust also ensure that no transaction is perennially aborted because it never has a sufficiently high priority. (Note that, in both schernes, the higher-priority transaction is never aborted.) When a transaction is aborted and restarted, it should be given the same timestamp it had originally. Reissuing timestarnps in this way ensures that each transaction will eventually becorne the oldest transaction, and therefore the one with the highest priority, and will get all the locks it requires.

The wait-die scheme is nonpreemptive; only a transaction requesting a lock can be aborted. As a transaction grows older (and its priority increases), it tends to wait for more and rnore younger transactions. A younger transaction that conflicts with an older transaction may be repeatedly aborted (a disadvantage with respect to wound-wait), but on the other hand, a transaction that has all the locks it needs is never aborted for deadlock reasons (an advantage with respect to wound-wait, which is preemptive).

A variant of 2PL, called Conservative 2PL, can also prevent deadlocks. Under Conservative 2PL, a transaction obtains all the locks it will ever need when it begins, or blocks waiting for these locks to become available. This scheme ensures that there will be no deadlocks, and, perhaps lllore important, that a transaction that already holds some locks will not block waiting for other locks. If lock contention is heavy, Conservative 2PL can reduce the time that locks are held on average, because transactions that hold locks are never blocked. The trade-off is that a transaction acquires locks earlier, and if lock contention is low, locks are held longer under Conservative 2PL. From a practical perspective, it is hard to know exactly what locks are needed ahead of time, and this approach leads to setting more locks than necessary. It also has higher overhead for setting locks because a transaction has to release all locks and try to obtain thern all over if it fails to obtain even one lock that it needs. This approach is therefore not used in practice.

17.5 SPECIALIZED LOCKING TECHNIQUES

Thus far we have treated a database as a *fixed* collection of *independent* data objects in our presentation of locking protocols. We now relax each of these restrictions and discuss the consequences.

If the collection of database objects is not fixed, but can grow and shrink through the insertion and deletion of objects, we must deal with a subtle cOlnplication known as the *phantom problem,* which was illustrated in Section 16.6.2. We discuss this problern in Section 17.5.1.

Although treating a database as an independent collection of objects is adequate for a discussion of serializability and recoverability, luuch better perfc)l'rnance can sOlnetilnes be obtained using protocols that recognize and exploit the relationships between objects. We discuss two such cases, nalnely, locking in tree-structured indexes (Section 17.5.2) and locking a collection of objects with containnent relationships between the1n (Section 17.5.3).

17.5.1 Dynamic Databases and the Phantom Problem

Consider the following exarnple: rrransaction *TI* scans the Sailors relation to find the oldest sailor for each of the *rating* levels 1 and 2. First, *TI* identifies and locks all pages (assurning that page-level locks are set) containing sailors with rating 1 and then finds the age of the oldest sailor, which is, say, 71. Next, transaction *T2* inserts a new sailor with rating 1 and age 96. Observe that this new Sailors record can be inserted onto a page that does not contain other sailors with rating 1; thus, an exclusive lock on this page does not conflict with any of the locks held by *TI*. *T2* also locks the page containing the oldest sailor with rating 2 and deletes this sailor (whose age is, say, 80). *T2* then comrnits and releases its locks. Finally, transaction *T1* identifies and locks pages containing (all remaining) sailors with rating 2 and finds the age of the oldest such sailor, which is, say, 63.

The result of the interleaved execution is that ages 71 and 63 are printed in response to the query. If *T1* had run first, then *T2*, we would have gotten the ages 71 and 80; if *T2* had run first, then *T*1, we would have gotten the ages 96 and 63. Thus, the result of the interleaved execution is not identical to any serial exection of *TI* and 1'2, even though both transactions follow Strict 2PL and cOlnmit. The problem is that *T1* assurnes that the pages it has locked include *all* pages containing Sailors records with rating 1, and this assurnption is violated when *rT2* inserts a new such sailor on a different page.

'rhe flaw is not in the Strict 2PL protocol. R,ather, it is in *T1*'s irnplicit assurnption that it has locked the set of all Sailors records with *rating* value 1. *T1*'s sernantics requires it to identify all such records, but locking pages that contain such records *at a given tirne* does not prevent new "phantom' records frorn being added on other pages. *T*1 has therefore *not* locked the set of desired Sailors records.

Strict 2PL guarantees conflict serializability; indeed, there are no cycles in the precedence graph for this exarnple because conflicts are defined with respect to objects (in this example, pages) read/written by the traJlsactions. However, because the set of objects that *should* have been locked by *TI* was altered by the actions ofT2, the olltcorne of the schedule differed frolll the outcorne of any

serial execution. This exalllple brings out an irnportant point about conflict serializability: If new itenls are added to the database, conflict serializability does not guarantee serializability.

A closer look at how a transaction identifies pages containing Sailors records with *rating* 1 suggests how the problenl can be handled:

- If there is no index and all pages in the file rnust be scanned, *T1* IUUSt someho\v ensure that no new pages are added to the file, in addition to locking all existing pages.
- If there is an index on the *rating* field, *T1* can obtain a lock on the index page—again, assurning that physical locking is done at the page level-that contains a data entry with *rating*= 1. If there are no such data entries, that is, no records with this *rating* value, the page that *would* contain a data entry for *rating=l* is locked to prevent such a record from being inserted. Any transaction that tries to insert a record with *rating=-l* into the Sailors relation 11lUSt insert a data entry pointing to the new record into this index page and is blocked until *T1* releases its locks. This technique is called index locking.

Both techniques effectively give *T1* a lock on the set of Sailors records with *rating=l:* Each existing record with *rating=l* is protected frolll changes by other transactions, and additionally, new records with *rating=l* cannot be inserted.

An independent issue is how transaction 7'1 can efficiently identify and lock the index page containing *rating*=1. We discuss this issue for the case of tree-structured indexes in Section 17.5.2.

\Ve note that index locking is a special case of a luore general concept called predicate locking. In our exalnple, the lock on the index page irnplieitly locked all Sailors records that satisfy the logical predicate *rating*= 1. 1V101'e generally, we can support irnplicit locking of all records that rnatch an arbitra,ry predicate. General predicate locking is expensive to irnplenlent and therefore not cOllullonly used.

17.5.2 Concurrency Control in B+ Trees

A straightforward approach to concurrency control for B+ trees and ISAM indexes is to ignore the index structure, treat each page as a data object, and use senne version of 2PL. This siInplistic locking strategy would lead to very high lock contention in the higher levels of the tree, because every tree search begins at the root and proceeds along sorne path to a leaf node. Fortunately, Innch Inore efficient locking protocols that exploit the hierarchical structure of a tree

index are known to reduce the locking overhead while ensuring seriaIizability and recoverability.We discuss sorne of these approaches briefly, concentrating on the search and insert operations.

Two observations provide the necessary insight:

1. The higher levels of the tree only direct searches. All the 'real' data is in the leaf levels (in the fornat of one of the three alternatives for data entries).

2. For inserts, a node must be locked (in exclusive rnode, of course) only if a split can propagate up to it frorn the modified leaf.

Searches should obtain shared locks on nodes, starting at the root and proceeding along a path to the desired leaf. The first observation suggests that a lock on a node can be released as soon as a lock on a child node is obtained, because searches never go back up the tree.

A conservative locking strategy for inserts would be to obtain exclusive locks on all nodes as we go down from the root to the leaf node to be modified, because splits can propagate all the way from a leaf to the root. However, once we lock the child of a node, the lock on the node is required only in the event that a split propagates back to it. In particular, if the child of this node (on the path to the modified leaf) is not full when it is locked, any split that propagates up to the child can be resolved at the child, and does not propagate further to the current node. Therefore, when we lock a child node, we can release the lock on the parent if the child is not full. The locks held thus by an insert force any other transaction following the sarne path to wait at the earliest point (i.e., the node nearest the root) that rnight be affected by the insert. The technique of locking a child node and (if possible) releasing the lock on the parent is called lock-coupling, or crabbing (think of how a crab walks, and cornpare it to how we proceed down a tree, alternately releasing a lock on a parent and setting a lock on a child).

We illustrate B+ tree locking using the tree in Figure 17.5. To search for data entry 38*, a transaction Ti rnust obtain an *S* lock on node *A*, read the contents and deterrnine that it needs to examine node *B*, obtain an *S* lock on node *B* and release the lock on *A*, then obtain an *S* lock on node C' and release the lock on *B*, then obtain an *S* lock on nodeD and release the lock on ().

Ti always rnaintains a lock all one node in the path, to force new transactions that want to read or nlodify nodes on the sarne path to wait until the current transaction is done. If transaction Tj wants to delete 38*, for exarnple, it rnust also traverse the path fro111 the root to node *D* and is forced to wait until Ti

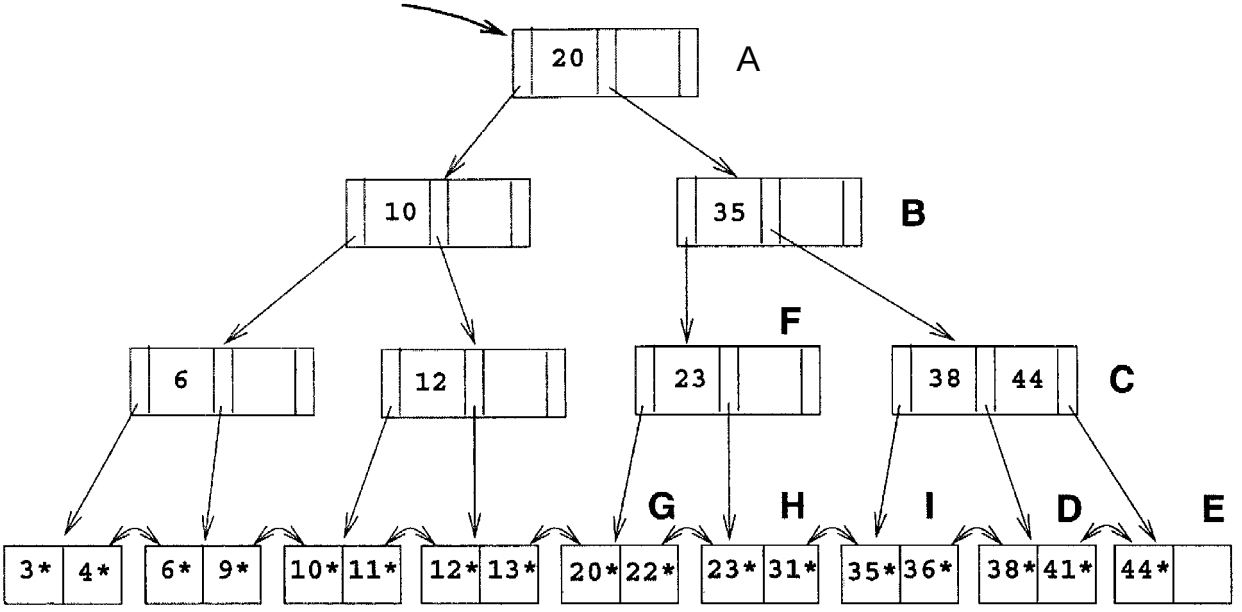


Figure 17.5 B+ Tree Locking Example

is done. Of course, if SOl11e transaction *Tk* holds a lock on, say, node C before *Ti* reaches this node, *Ti* is similarly forced to wait for *Tk* to complete.

To insert data entry 45*, a transaction 111USt obtain an *S* lock on node *A*, obtain an *S* lock on node *B* and release the lock on *A*, then obtain an *S* lock on node C (observe that the lock on *B* is *not* released, because C is full), then obtain an X lock on node *E* and release the locks on C and then *B*. Because node *E* has space for the new entry, the insert is accomplished by modifying this node.

In contrast, consider the insertion of data entry 25*. Proceeding as for the insert of 45*, we obtain an X lock on node *H*. Unfortunately, this node is full and must be split. Splitting *H* requires that we also rnodify the parent, node *F*, but the transaction has only an *S* lock on *F*. Thus, it must request an upgrade of this lock to an X lock. If no other tra,nsaction holds an *S* lock on *F*, the upgrade is granted, and since *F* has space, the split does not propagate further and the insertion of 25* can proceed (by splitting *If* and locking G to modify the sibling pointer in *I* to point to the newly created node). However, if another transaction holds an *S* lock on node *F*, the first transaction is suspended until this transaction releases its Slack.

Observe that if another transaction holds an *S* lock on *F* and also wants to access node *H*, we have a deadlock because the first transaction has an X lock on *If*. The preceding exarnple also illustrates an interesting point about sibling pointers: When we split leaf node *H*, the new node *must* be added to the *left* of *Ii*, since otherwise the node whose sibling pointer is to be changed would be node 1, which has a different parent. To rnodify a sibling pointer on *I*, we

'would have to lock its parent, node C (and possibly ancestors of C, in order to lock C).

Except for the locks on intermediate nodes that we indicated could be released early, senne variant of 2PL HUlst be used to govern when locks can be released, to ensure serializability and recoverability.

T'his approach irllproves considerably on the naive use of 2PL, but several exclusive locks are still set unnecessarily and, although they are quickly released, affect perfon.nance substantially. One way to iInprove perforlllance is for inserts to obtain shared locks instead of exclusive locks, except for the leaf, which is locked in exclusive 11lode. In the vast rnajority of cases, a split is not required and this approach works very well. If the leaf is full, however, we Blust upgrade from shared locks to exclusive locks for all nodes to which the split propagates. Note that such lock upgrade requests can also lead to deadlocks.

The tree locking ideas that we describe illustrate the potential for efficient locking protocols in this very important special case, but they are not the current state of the art. The interested reader should pursue the leads in the bibliography.

17.5.3 Multiple-Granularity Locking

Another specialized locking strategy, called **multiple-granularity locking,** allows us to efficiently set locks on objects that contain other objects.

For instance, a database contains several files, a file is a collection of pages, and a page is a collection of records. A transaction that expects to access rnost of the pages in a file should probably set a lock on the entire file, rather than locking individual pages (or records) when it needs thern. Doing so reduces the locking overhead considerably. On the other hand, other tra,nsactions that require access to parts of the file---even parts not needed by this transaction---are blocked. If a transaction accesses relatively few pages of the file, it is better to lock only those pages. Sirnilarly, if a transaction accesses several records on a page, it should lock the entire page, and if it accesses just a few records, it should lock just those records.

The question to be addressed is how a lock rnanager can efficiently ensure that a page, for exaruple, is not locked by a transaction while another transaction holds a conflicting lock on the file containing the page (a.nd therefore, irnplicitly, on the page).

The idea is to exploit the hierarchical nature of the 'contains' relationship. A database contains a set of files, each file contains a set of pages, and each page contains a set of records. This containunent hierarchy can be thought of as a tree of objects, \vhere each node contains all its children. (The approach can easily be extended to cover hierarchies that are not trees, but we do not discuss this extension.) A lock on a node locks that node and, irnplicitly, all its descendants. (Note that this interpretation of a lock is very different fron1 B+ tree locking, where locking a node does *not* lock any descendants ilnplicitly.)

In addition to shared (8) and exclusive *(XO)* locks, rnultiple-granularity locking protocols also use two new kinds of locks, called intention shared (18) and intention exclusive *(IX)* locks. 18 locks conflict only with X locks. IX locks conflict with 8 and X locks. To lock a node in *S* (respectively, X) luode, a transaction must first lock all its ancestors in 18 (respectively, 1*X)* rllode. Thus, if a transaction locks a node in 8 rIlode, no other transaction can have locked any ancestor in X rnode; siInilarly, if a transaction locks a node in X mode, no other transaction can have locked any ancestor in 8 or X mode. This ensures that no other transaction holds a lock on an ancestor that conflicts with the requested 8 or X lock on the node.

A common situation is that a transaction needs to read an entire file and modify a few of the records in it; that is, it needs an 8 lock on the file and an 1X lock so that it can subsequently lock sorne of the contained objects in X mode. It is useful to define a new kind of lock, called an 81*X* lock, that is logically equivalent to holding an 8 lock and an *IX* lock. A transaction can obtain a single 81*X* lock (which conflicts with any lock that conflicts with either *S* or *IX)* instead of an 8 lock and an *IX* lock.

A subtle point is that locks rnust be released in leaf-to-root order for this protocol to work correctly. To see this, consider what happens when a transaction *Ti* locks all nodes on a path frolH the root (corresponding to the entire database) to the node corresponding to sorne page *p* in 18 rnode, locks *p* in *S* rHode, and then releases the lock on the root node. Another transaction *T*j could now obtain an *X* lock on the root. This lock iInplicitly gives *Tj* an *X* lock on page *p,* which conflicts with the 8 lock currently held by *Ti*.

Multiple-granularity locking lllust be used with 2PL to ensure serializability. The 2PL protocol dictates when locks can be released. At that tirne, locks obtained using rIlultiple-granularity locking can be released and IllUSt be released in leaf-to-root order.

Finally, there is the question of how to decide what granularity of locking is appropriate for a given transaction. One approach is to begin by obtaining fine granularity locks (e.g., at the record level) and, after the transaction requests

Lock Granularity: SOfie database systeIlls allow programmers to override the default mechanisllt for choosing a lock granularity. For exalnple, Microsoft SQL Server allows users to select page locking instead of table locking, using the keyword PAGLOCK. IBM's DB2 UDB allows for explicit table-level locking.

a certain nUlnber of locks at that granularity, to start obtaining locks at the next higher granularity (e.g., at the page level). This procedure is called **lock** escalation.

17.6 CONCURRENCY CONTROL WITHOUT LOCKING

Locking is the most widely used approach to concurrency control in a DBMS, but it is not the only one. We now consider some alternative approaches.

17.6.1 Optimistic Concurrency Control

Locking protocols take a pessimistic approach to conflicts between transactions and use either transaction abort or blocking to resolve conflicts. In a systenl with relatively light contention for data objects, the overhead of obtaining locks and following a locking protocol must nonetheless be paid.

In optimistic concurrency control, the basic premise is that most transactions do not conflict with other transactions, and the idea is to be as permissive as possible in allowing transactions to execute. Transactions proceed in three phases:

1. Read: The transaction executes, reading values froIn the database and writing to a private workspace.
2. Validation: If the transaction decides that it wants to c0l111uit, the DBIvIS checks whether the transaction could possibly have conflicted with any other concurrently executing transaction. If there is a possible conflict, the transaction is aborted; its private workspace is cleared and it is restarted.
3. Write: If validation deterrnines that there are no possible confliets, the changes to data objects 111ade by the transaction in its private workspace are copied into the database.

If, indeed, there are few confiicts, and validation can be done efficiently, this approach should lead to better performance than locking. If there are rnany

conflicts, the cost of repeatedly restarting transactions (thereby wasting the work they've done) hurts perfornlance significantly.

Each transaction *Ti* is assigned a thnestamp $TS(Ti)$ at the beginning of its validation phase, and the validation criterion checks whether the tiITlestalnp-ordering of transactions is an equivalent serial order. For every pair of transactions *Ti* and *Tj* such that $TS(Ti) < TS(Tj)$, one of the following validation conditions ITIUSt hold:

1. *Ti* completes (all three phases) before *Tj* begins.
2. *Ti* completes before *Tj* starts its Write phase, and *Ti* does not write any database object read by *Tj.*
3. *Ti* completes its Read phase before *Tj* completes its Read phase, and *Ti* does not write any database object that is either read or written by *Tj.*

To validate *Tj,* we must check to see that one of these conditions holds with respect to each comlnitted transaction *Ti* such that $TS(Ti) < TS(Tj)$. Each of these conditions ensures that *Tj's* modifications are not visible to *Ti.*

Further, the first condition allows *Tj* to see some of *Ti's* changes, but clearly, they execute completely in serial order with respect to each other. The second condition allows *Tj* to read objects while *Ti* is still modifying objects, but there is no conflict because *Tj* does not read any object rnodified by Ti. Although *Tj* might overwrite some objects written by *Ti,* all of *Ti's* writes precede all of *Tj's* writes. The third condition allows *Ti* and *Tj* to write objects at the same time and thus have even IT10re overlap in time than the second condition, but the sets of objects written by the two transactions cannot overlap. Thus, no RW, WR, or WW conflicts are possible if any of these three conditions is met.

Checking these validation criteria requires us to maintain lists of objects read and written by each transaction. Further, while one transaction is being validated, no other transaction can be allowed to commit; otherwise, the validation of the first transaction might miss conflicts with respect to the newly committed transaction. The Write phase of a validated transaction rnust also be completed (so that its effects are visible outside its private workspace) before other transactions can be validated.

A synchronization rnechanisrn such as a critical section can be used to ensure that at most one transaction is in its (colllbined) Validation/Write phases at any tirne. (When a process is executing a critical section in its code, the systern suspends all other processes.) Obviously, it is irnportant to keep these phases as short HS possible in order to rniniruize the irnpact on concurrency. If copies of rnodified objects have to be copied frorn the private workspace, this

can rnake the \Vrite phase long. An alternative approach (which carries the penalty of poor physical locality of objects, such as B·+· tree leaf pages, that rnust be clustered) is to use a level of indirection. In this schernc, every object is accessed via a logical pointer, and in the Write phase, we sirnply switch the logical pointer to point to the version of the object in the private workspace, instead of copying the 0bject.

Clearly, it is not the case that optiInistic concurrency control has no overheads; rather, the locking overheads of lock-based approaches are replaced with the overheads of recording read-lists and write-lists for transactions, checking for conflicts, and copying changes frorn the private workspace. Sirnilarly, the irnplicit cost of blocking in a lock-based approach is replaced by the implicit cost of the work wasted by restarted transactions.

Improved Conflict Resolution[1]

Optirnistic Concurrency Control using the three validation conditions described earlier is often overly conservative and unnecessarily aborts and restarts transactions. In particular, according to the validation conditions, Ti cannot write any object read by Tj. IIowever, since the validation is airned at ensuring that Ti logically executes before Tj, there is no harm if Ti writes all data items required by Tj before Tj reads theIn.

The problerIl arises because we have no way to tell when Ti wrote the object (relative to Tj's reading it) at the tirne we validate Tj, since all we have is the list of objects written by Ti and the list read by Tj. Such false conflicts can be alleviated by a finer-grain resolution of data conflicts, using rnechanisrI1s very sinlilar to locking.

The basic idea is that each transaction in the Read pha.se tells the DBMS about iteIIls it is reading, and ·when a transaction Ti is cornrnitted (and its writes are accepted), the DBMS checks whether any of the iterns written by Ti are being read by any (yet to be validated) transaction Tj. If so, we know thatTj's validation rnust eventually fail. We can either allow Ti to discover this when it is validated (the die policy) or kill it and restart it innnediately (the kill policy).

The details are as follo\vs. Before reading a data iterrl, (1 transaction Tenters an access entry in a hash table. The access entry contains the *transaction id,* a *data object id,* and a *rn..odified* flag (initially set to false), and entries are hashed on the data object id. A terl1porary exclusive lock is obtained on the

[1]We thank Alexander Thomasian for writing this section.

hash bucket containing the entry, and the lock is held \vhile the read data iteIll is copied frolll the database buffer into the private 'workspace of the transactioll.

During validation of T the hash buckets of all data objects accessed by T are again locked (in exclusive 11lode) to check if T has encountered any data conflicts. T has encountered a conflict if the *rnodified* flag is set to true in one of its access entries. (This assumes that the 'die' policy is being used; if the 'kill' policy is used, T is restarted when the flag is set to true.)

If T is successfully validated, we lock the hash bucket of each object lnodified by T, retrieve all access entries for this object, set the *rnodified* flag to true, and release the lock on the bucket. If the 'kill' policy is used, the transactions that entered these access entries are restarted. We then complete T's Write phase.

It seems that the 'kill' policy is always better than the 'die' policy, because it reduces the overall response time and wasted processing. However, executing T to the end has the advantage that all of the data items required for its execution are prefetched into the database buffer, and restarted executions of T will not require disk I/O for reads. This assumes that the database buffer is large enough that prefetched pages are not replaced, and, 1nore irnportant, that access invariance prevails; that is, successive executions of T require the same data for execution. When T is restarted its execution tirne is nluch shorter than before because no disk I/O is required, and thus its chances of validation are higher. (Of course, if a transaction has already completed its Read phase once, subsequent conflicts should be handled using the 'kill' policy because all its data objects are already in the buffer pool.)

17.6.2 Timestamp-Based Concurrency Control

In lock-based concurrency control, conflicting actions of different transactions are ordered by the order in which locks are obtained, and the lock protocol extends this ordering on actions to transactions, thereby ensuring serializability. In optirrlistic concurrency control, a timestamp ordering is irnposed on transactions and validation checks that all conflicting actions occurred in the saIne order.

Tinlcstarnps can also be used in another \vay: Each transaction can be assigned a tirnestanlp at startup, and we can ensure, at execution tirne, that if action ai of transaction Ti conflicts \vith action aj of transaction Tj, ai occurs before aj if $TS(Ti) < TS(Tj)$. If an action violates this ordering, the transaction is aborted and restarted.

To irnplernent this concurrency control scheme, every database object 0 is given a read tirnestampRTS(0) and a write timestamp $WTS(0)$. If transaction T wants to read object 0, and $TS(T) < WTS(O)$, the order of this read with respect to the most recent write on 0 would violate the timestamp order between this transaction and the writer. Therefore, T is aborted and restarted *with a new, larger timestarnp.* If $TS(T) > WTS(O)$, Treads 0, and $RTS(O)$ is set to the larger of $RTS(O)$ and $TS(T)$. (Note that a physical change—the change to $RTS(O)$-is written to disk and recorded in the log for recovery purposes, even on reads. This write operation is a significant overhead.)

Observe that if T is restarted with the same timestamp, it is guaranteed to be aborted again, due to the saIne conflict. Contrast this behavior with the use of timestamps in 2PL for deadlock prevention, where transactions are restarted with the *same* timestarnp as before to avoid repeated restarts. This shows that the two uses of timestamps are quite different and should not be confused.

Next, consider what happens when transaction T wants to write object 0:

1. If $TS(T) < RTS(O)$, the write action conflicts with the most recent read action of 0, and T is therefore aborted and restarted.
2. If $TS(T) < WTS(O)$, a naive approach would be to abort T because its write action conflicts with the most recent write of 0 and is out of timestamp order. However, we can safely ignore such writes and continue. Ignoring outdated writes is called the Thomas Write Rule.
3. Otherwise, T writes 0 and $WTS(O)$ is set to $TS(T)$.

The Thomas Write Rule

We now consider the justification for the TholIlas Write Rule. If $TS(T) < WTS(O)$, the current write action has, in effect, been made obsolete by the rnost recent write of 0, which *follows* the current write according to the tirnestalnp ordering. We can think of *T's* write action as if it had occurred irnrnediately *before* the rnost recent write of 0 and was never read by anyone.

If the Thomas vVrite Rule is not used, that is, T is aborted in case (2), the tirnestamp protocol, like 2PL, allows only conflict serializable schedules. If the Tho1l1aS Write R,ule is used, some schedules are perrnitted that are not conflict serializable, as illustrated by the schedule in Figure 17.6.[2] Because *T2's* write follows Tl's read and precedes *Tl's* write of the sanle object, this schedule is not conflict serializable.

21n the other direction, 2PL pennits some schedules that are not allowed by the timestamp algorithm with the Thomas Write Rule; see Exercise 17.7.

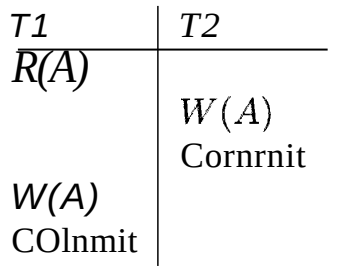


Figure 17.6 A Serializable Schedule 'rhat Is Not Conflict Serializable

The Thomas Write Rule relies on the observation that $T2$'s write is never seen by any transaction and the schedule in Figure 17.6 is therefore equivalent to the serializable schedule obtained by deleting this write action, which is shown in Figure 17.7.

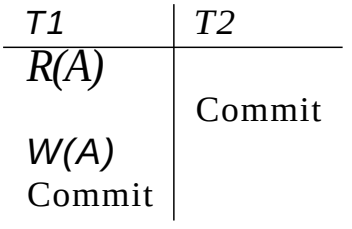


Figure 17.7 A Conflict Serializable Schedule

Recoverability

Unfortunately, the timestamp protocol just presented permits schedules that are not recoverable, as illustrated by the schedule in Figure 17.8. If $TS(T1) = 1$ and $T8(T2) = 2$, this schedule is permitted by the timestalnp protocol (with or without the 1"ho111as Write Rule). The tiInestalnp protocol can be modified to disallow such schedules by buffering all write actions until the transaction COIDlnits. In the example, when TI wants to write A, $WTS(A)$ is updated to reflect this action, but the change to A. is not carried out irrllnediately; instead, it is recorded in a private workspace, or buffer. When $T2$ wants to read A subsequently, its thnestamp is cornpared with $I¥TS(A)$, and the read is seen to be perrnissible. However, $T2$ is blocked until $T1$ cornpletes. If $T1$ cornrnits, its change to A is copied fro111 the buffer; other\vise, the changes in the buffer are discarded. $T2$ is then allowed to read A.

This blocking of $T2$ is sinlilar to the effect of $T1$ obtaining an exclusive lock on A. Nonetheles8, even with this modification, the tirnestarnp protocol perrnits sorne schedules not perrnitted by 2PL; the two protocols are not quite the same. (See Exercise 17.7.)

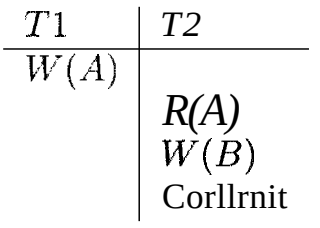


Figure 17.8 An Unrecoverable Schedule

Because recoverability is essential, such a modification must be used for the timestamp protocol to be practical. Given the added overhead this entails, on top of the (considerable) cost of maintaining read and write tilnestamps, thnestamp concurrency control is unlikely to beat lock-based protocols in centralized systems. Indeed, it has been used mainly in the context of distributed database systems (Chapter 22).

17.6.3 Multiversion Concurrency Control

This protocol represents yet another way of using timestamps, assigned at startup time, to achieve serializability. The goal is to ensure that a transaction never has to wait to read a database object, and the idea is to maintain several versions of each database object, each with a write timestamp, and let transaction *Ti* read the most recent version whose timestarnp precedes *TS(Ti).*

If transaction 1'i wants to write an object, we must ensure that the object has not already been read by sonle other transaction *Tj* such that *TS(Ti)* < *1'S(Tj).* If we allow *Ti* to write such an object, its change should be seen by *Tj* for serializability, but obviously *Tj,* which read the object at SaIne tinle in the past, will not see *Ti*'s change.

To check this condition, every object also has an associated read timestarnp, and whenever a transaction reads the object, the read timestamp is set to the maxhuuru of the current read tilnestarnp and the reader's tirnestarnp. If *7',t* wants to write an object 0 and *TS(Ti)* < *RTS(O), Ti* is aborted and restarted with a new, larger timestamp. Otherwise, *Ti* creates a new version of 0 and sets the read and write tirnestarnps of the new version to *7'S(Ti).*

The drawbacks of this sehenle are similar to those of tirnestarnp concurrency control, and in addition, there is the cost of rnaintaining versions. On the other hand, reads are never blocked, which can be irnportant for workloads dorninated by transactions that only read values frorn the database.

What Do Real Systems Do? IBM DB2, Informix, Microsoft SQL Server, and Sybase ABE use Strict 2PL or variants (if a transaction requests a lower than SERIALIZABLE SQL isolation level; see Section 16.6). IVlicrosoft SQL Server also supports rnodifieation timestamps so that a transaction can run without setting locks and validate itself (do-it-yourself OptirnisticConC1:1rrency Control!). Oracle 8 uses a lllultiversion concurrency control scherne in which readers never wait; in fact, readers never get locks and detect conflicts by checking if a block changed since they read it. All these systerlls support rnultiple-granu1arity locking, with support for table, page, and row level locks. All deal with deadlocks using waits-for graphs. Sybase ASIQ supports only table-level locks and aborts a transaction if a lock request fails-·--updates (and therefore conflicts) are rare in a data warehouse, and this simple scheme suffices.

17.7 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- When are two schedules *conflict equivalent*? What is a *conflict serializable* schedule? What is a *strict* schedule? (Section 17.1)
- What is a *precedence graph* or *serializability graph?* Ilow is it related to conflict serializability? How is it related to two-phase locking? (Section 17.1)
- What does the *lock manager* do? Describe the *lock table* and *transaction table* data structures and their role in lock management. (Section 17.2)
- Discuss the relative merits of *lock upgrades* and *lock downgrades.* (Section 17.3)
- Describe and cornpare deadlock detection a,nd deadlock prevention schernes. Why are detection schernes rnore cornrnonly used? (Section 17.4)
- If the collection of database objects is not fixed, but can gro\v and shrink through insertion and deletion of objects, we lnust deal with a subtle corllplication known as the *phantorn problern.* Describe this problern and the index locking approach to solving the probleln. (Section 17.5.1)
- In tree index structures, locking higher levels of the tree can becorne a perforrnanee bottleneck. Explain why. I)escribe specialized locking techniques that address the problenl, and explain why they work correctly despite not lJeing two-phase. (Section 17.5.2)
- *Multiple-granularity locki'ng* enables us to set locks on objects that contain other objects, thus implicitly locking all contained objects. Why is this approach irnportant and how does it work? (Section 17.5.3)

- In *optimistic concurrency control*, no locks are set and transactions read and rnodify data objects in a private workspace. How are conflicts between transactions detected and resolved in this approach? (Section 17.6.1)
- In *tirnestamp-based concurrency control*, transactions are assigned a timestarnp at startup; how is it used to ensure serializability? How does the *Thomas Write Rule* improve concurrency? (Section 17.6.2)
- Explain why tinlestamp-based concurrency control allows schedules that are not recoverable. Describe how it can be modified through *buffering* to disallow such schedules. (Section 17.6.2)
- Describe *multiversion concurrency control.* What are its benefits and disadvantages in comparison to locking? (Section 17.6.3)

EXERCISES

Exercise 17.1 Answer the following questions:

1. Describe how a typical lock manager is implemented. Why must lock and unlock be atomic operations? What is the difference between a lock and a *latch?* What are *convoys* and how should a lock manager handle them?
2. Compare *lock downgrades* with upgrades. Explain why downgrades violate 2PL but are nonetheless acceptable. Discuss the use of *update* locks in conjunction with lock downgrades.
3. Contrast the timestamps assigned to restarted transactions when tinwstanlps are used for deadlock prevention versus when timestamps are used for concurrency control.
4. State and justify the Thomas Write Rule.
5. Show that, if two schedules are conflict equivalent, then they are view equivalent.
6. Give an example of a serializable schedule that is not strict.
7. Give an example of a strict schedule that is not serialiable.
8. Motivate and describe the use of locks for improved conflict resolution in Optinlistic Concurrency Control.

Exercise 17.2 Consider the following classes of schedules: *serializable, confiict-serializable, view-serializable, recoverable, avoids-cascading-aborts,* and *strict.* For each of the following schedules, state which of the preceding classes it belongs to. If you cannot decide whether a schedule belongs in a certain class based on the listed actions, explain briefly.

The actions are listed in the order they are scheduled and prefixed with the transaction name. If a commit or abort is not shown, the schedule is incomplete; assurne that abort or cornrnit lllust follow all the listed actions.

1. Tl:R(X), T2:R(X), Tl:W(X), T2:W(X)
2. Tl:W(X), T2:R(Y), Tl:R(Y), T2:R(X)

3. Tl:R(X), T2:R(Y), T3:W(X), T2:R(X), Tl:R(Y)
4. Tl:R(X), T1:R(Y), T1:W(X), T2:R(Y), T3:W(Y), Tl:W(X), T2:R(Y)
5. Tl:R(X), T2:W(X), Tl:W(X), T2:Abort, Tl:Cmnmit
6. Tl:R(X), T2:W(X), T1:W(X), T2:Comrnit, Tl:Comm.it
7. T1:W(X), T2:R(X), 1'l:W(X), 1'2:Abort, T.1:COlllllllit
8. Tl:W(X), T2:R(X), Tl:W(X), T2:Conunit, Tl:Col1unit
9. Tl:W(X), T2:R(X), Tl:W(X), T2:Commit, Tl:Abort
10. 1'2: R(X), 1'3:W(X), T3:Cmnrnit, Tl:W(Y), Tl:Commit, T2:R(Y), T2:W(Z), T2:Colllmit
11. Tl:R(X), T2:W(X), T2:Cornrnit, Tl:W(X), Tl:Colllmit, T:3:R(X), T3:Collnnit
12. Tl:R(X), T2:W(X), Tl:W(X), T3:R(X), Tl:Comllllit, T2:Corn111it, 1'3:Comlnit

Exercise 17.3 Consider the following concurrency control protocols: 2PL, Strict 2PL, Conservative 2PL, Optimistic, Tilnestamp without the Thomas Write Rule, 1'ilnestamp with the Thomas Write Rule, and Multiversion. For each of the schedules in Exercise 17.2, state which of these protocols allows it, that is, allows the actions to occur in exactly the order shown.

For the timestamp-based protocols, assurne that the timestamp for transaction *Ti* is i and that a version of the protocol that ensures recoverability is used. Further, if the Thomas Write Rule is used, show the equivalent serial schedule.

Exercise 17.4 Consider the following sequences of actions, listed in the order they are submitted to the DBMS:

- Sequence 81: Tl:R(X), T2:W(X), T2:W(Y), T3:W(Y), Tl:W(Y), Tl:Commit, T2:Commit, T3:Commit
- Sequence 82: Tl:R(X), T2:W(Y), T2:W(X), T3:W(Y), Tl:W(Y), Tl:C0111mit, T2:Commit, T3:Commit

For each sequence and for each of the following concurrency control rnechanislns, describe how the concurrency control mechanislll handles the sequence.

Assurne that the tirnestarnp of transaction *Ti* is *i*. Fbr lock-based concurrency control rnechaniS111S, add lock and unlock requests to the previous sequence of actions as per the locking protocol. The DBMS processes actions in the order shown. If a transaction is blocked, assume that all its actions are queued until it is reslllned; the DBMS continues with the next action (according to the listed sequence) of an unblocked transaction.

1. Strict 2PL with tiluestamps used for deadlock prevention.
2. Strict 2PL with deadlock detection. (Show the waits-for graph in case of deadlock.)
3. Conservative (and Strict, i.e., with locks held until end-of-transaction) 2PL.
4. Optimistic concurrency control.
5. Tiruestarup concurrency control with buffering of reads and writes (to ensure recoverability) and the Tholnas Write Rule.
6. rvluitiversioll concurrency control.

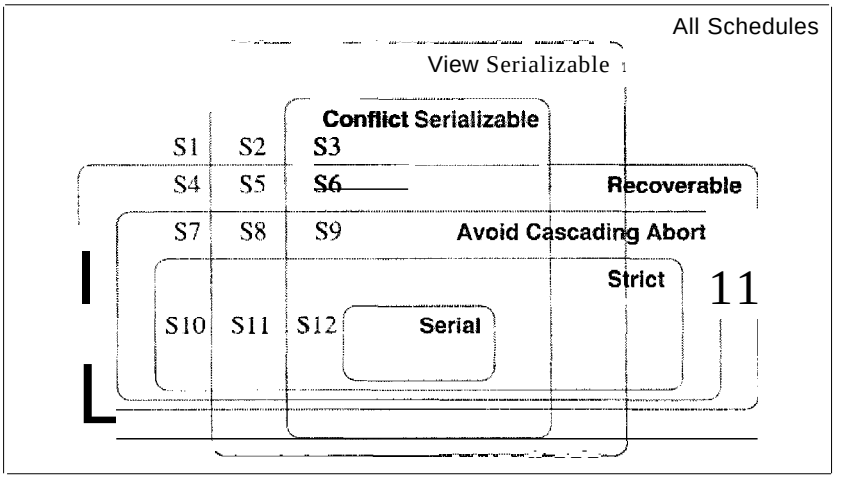


Figure 17.9 Venn Diagram for Classes of Schedules

Exercise 17.5 For each of the following locking protocols, assulning that every transaction follows that locking protocol, state which of these desirable properties are ensured: serializability, conflict-·serializability, recoverability, avoidance of cascading aborts.

1. Always obtain an exclusive lock before writing; hold exclusive locks until end-of-transaction. No shared locks are ever obtained.
2. In addition to (1), obtain a shared lock before reading; shared locks can be released at any time.
3. As in (2), and in addition, locking is two-phase.
4. As in (2), and in addition, all locks held until end-of-transaction.

Exercise 17.6 The Venn diagranl (frorn [76]) in Figure 17.9 shows the inclusions between several classes of schedules. Give one exaulple schedule for each of the regions S1 through *S12* in the diagrarn.

Exercise 17.7 Briefly answer the following questions:

1. Draw a Venn diagram that ShO\V8 the inclusions between the classes of schedules perulittccl by the following concurrency control protocols: *2PL, Strict 2PL, Conservative 2PL, Optimistic, Timestamp without the Thomas Write Rule, Timestamp with the Thomas Write Rule,* and *Multiversion.*
2. Give one example schedule for each region in the diagrarIl.
3. Extend the Venn diagranl to include serializable and conflict-serializable schedules.

Exercise 17.8 Answer each of the follmving questions briefly. The questions are based on the following relational schern.a:

Ernp(*eid:* integer, *ename:* string, *age:* integer, *salary:* real, *did:* integer)
Dept(*did:* integer, *dname:* string, *flooT:* integer)

and on the fc)llowing update cornrnand:

replace (salary = 1.1 * EMP.salary) where EMP.ename = 'Santa'

1. Give an example of a query that would conflict with this comrnand (in a concurrency control sense) if both were run at the same tim.e. Explain what could go wrong, and how locking tuples would solve the probleIll.
2. Give an exarnple of a query or a cOHlInand that would conflict with this cOIl1ruand, such that the conflict could not be resolved by just locking individual tuples or pages but requires index locking.
3. Explain what index locking is and how it resolves the preceding conflict.

Exercise 17.9 SQL supports four isolation-levels and two access-rllodes, for a total of eight cornbinations of isolation-level and access-rnode. Each corubinatioll inlplicitly defines a class of transactions; the follO\ving questions refer to these eight classes:

1. For each of the eight classes, describe a locking protocol that allows only transactions in this class. Does the locking protocol for a given class make any assurnptiolls about the locking protocols used for other classes? Explain briefly.
2. Consider a schedule generated by the execution of several SQL transactions. Is it guaranteed to be conflict-serializable? to be serializable'? to be recoverable?
3. Consider a schedule generated by the execution of several SQL transactions, each of which has READ ONLY access-mode. Is it guaranteed to be conflict-serializable? to be serializable? to be recoverable?
4. Consider a schedule generated by the execution of several SQL transactions, each of which has SERIALIZABLE isolation-level. Is it guaranteed to be conflict-serializable? to be serializable? to be recoverable?
5. Can you think of a tinlCstarup-based concurrency control scheme that can support the eight classes of SQL transactions?

Exercise 17.10 Consider the tree shown In Figure 19.5. Describe the steps involved in executing each of the following operations according to the tree-index concurrency control algorithm discussed in Section 19.3.2, in terms of the order in which nodes are locked, unlocked, read, and written. Be specific about the kind of lock obtained and answer each part independently of the others, always starting with the tree shown in Figure 19.5.

1. Search for data entry 40*.
2. Search for all data entries k* with $k \leq 40$.
3. Insert data entry 62*.
4. Insert data entry 40*.
5. Insert data entries 62* and 75*.

Exercise 17.11 Consider a database organized in tenns of the following hierarachy of objects: The database itself is an object *(D),* and it contains two files *(FI* and *F'2),* each of which contains 1Q00 pages *(PI* ... PlOOO ancl *P*1001 ... P2000, respectively). Each page contains 100 records, and records are identified as $p : i$, where p is the page identifier and i is the slot of the record on that page.

I'vlultiple-granularity locking is used, with S, X, IS, IX and SIX locks, and database-level, file-level, page-level and record-level locking. For each of the following operations, indicate the sequence of lock requests that 1nust be generated by a transaction that wants to carry out (just) these operations:

1. Read record P1200 : 5.
2. Read records $P1200 : 98$ through P1205 : 2.
3. Read all (records on all) pages in file $F1$.
4. 'Read pages P500 through P520.
5. Read pages PIO through P980.
6. Read all pages in *PI* and (based on the values read) rnodify 10 pages.
7. Delete record P1200 : 98. (This is a blind write.)
8. Delete the first record frorn each page. (Again, these are blind writes.)
9. Delete all records.

Exercise 17.12 Suppose that we have only two types of transactions, T1 and *T2*. Transactions preserve database consistency when run individually. We have defined several *integrity constraints* such that the DBMS never executes any SQL statenwnt that brings the database into an inconsistent state. Assunle that the DBlVIS does not perform *any* concurrency control. Give an exarllple schedule of two transactions T1 and T2 that satisfies all these conditions, yet produces a database instance that is not the result of any serial execution of $T1$ and T2.

BIBLIOGRAPHIC **NOTES**

Concurrent access to B trees is considered in several papers, including [70,456,472,505,678]. Concurrency control techniques for Linear Hashing are presented in [240] and [543]. Multiple-granularity locking is introduced in [3:36] and studied further in [127, 449].

A concurrency control method that works with the ARIES recovery rnethod is presented in [545]. Another paper that considers concurrency control issues in the context of recovery is [492]. AlgorithrIls for building indexes without stopping the DBMS are presented in [548] and [9]. The performance of B tree concurrency control algorithnls is studied in [704]. Perforrnlance of various concurrency control algorithms is discussed in [16, 729, 735]. A good survey of concurrency control rnethods and their perforrnlance is [734]. [455] is a comprehensive collection of papers on this topic.

TilYwstarIlp-based multiversion concurrency control is studied in [6201. IvIultiversion concur--rency control algoritlllIIls are studied forrnally in [87J. Lock-based rnultiversion techniques are considered in [460]. Optirnistic concurrency control is introduced in [457]. The use of access invariance to irnprove conflict resolution in high-contention environrnents is discussed in [281] and [280]. Transaction management issues for real-tirne database systerns are discussed in [1, 15, 368, :382, 386, 448]. There is a large body of theoretical results on database concurrency control; [582, 89] offer thorough textbook presentations of this rnaterial.

18

CRASH RECOVERY

- What steps are taken in the ARIES method to recover fronl a DBMS crash?
- How is the log rnaintained during nonnal operation?
- How is the log used to recover frorn a crash?
- What infonnation in addition to the log is used during recovery?
- What is a checkpoint and why is it used?
- W'hat happens if repeated crashes occur during recovery?
- How is media failure handled?
- How does the recovery algorithnl interact with concurrency control?
- Key concepts: steps in recovery, analysis, redo, undo; ARIES, repeating history; log, LSN, forcing pages, WAL; types of log records, update, cornrnit, abort, end, cOlnpensation; transaction ta-· ble, lastLSN; dirty page table, recLSN; checkpoint, fuzzy checkpointing, master log record; rnedia recovery; interaction with concurrency control; shadow paging

Hurnpty Durnpty sat on a \vall.
lIurnpty Durnpty had a great fall.
A.ll the King's horses and all the King's tnen
Could not put lIllrnpty together again.

—Old nursery rhyrne

The recovery manager of a DBMS is responsible for ensuring two irnportant properties of transactions: Atornicity and durability. It ensures *atomicity* by undoing the actions of transactions that do not conlIllit and *durability* by rnaking sure that all actions of conunitted transactions survive system crashes (e.g., a core durnp caused by a bus error) and Inedia failures (e.g., a disk is corrupted).

1'"he recovery rnanager is one of the hardest cOlllponents of a DBMS to design and inlplernent. It rnust deal 'with a wide va,riety of database states because it is called on during systenl failures. In this chapter, we present the ARIES recovery algorithnl, which is conceptually sinlple, works well with a wide range of concurrency control rnechanisrns, and is being used in an increasing number of database syterns.

We begin with an introduction to ARIES in Section 18.1. We discuss the log, which a central data structure in recovery, in Section 18.2, and other recovery-related data structures in Section 18.3. We complete our coverage of recovery-related activity during normal processing by presenting the Write-Ahead Logging protocol in Section 18.4, and checkpointing in Section 18.5.

We discuss recovery frorn a crash in Section 18.6. Aborting (or rolling back) a single transaction is a special case of Undo, discussed in Section 18.6.3. We discuss media failures in Section 18.7, and conclude in Section 18.8 with a discussion of the interaction of concurrency control and recovery and other approaches to recovery. In this chapter, we consider recovery only in a centralized DBMS; recovery in a distributed DBMS is discussed in Chapter 22.

18.1 INTRODUCTION TO ARIES

ARIES is a recovery algorithrn designed to work with a steal, no-force approach. When the recovery rnanager is invoked after a crash, restart proceeds in three phases:

1. Analysis: Identifies dirty pages in the buffer pool (i.e., changes that have not been written to disk) and active transactions at the tiITle of the crash.
2. Redo: Repeats all actions, starting frOID an appropriate point in the log, and restores the database state to what it was at the tirne of the e1'a8h.
3. lJndo: Undoes the actions of transactions that did not cOllunit, so that the database reflects only the actions of cornrnitted transactions.

Consider the sirnple execution history illustrated in Figure 18.1. When the systeIIl is restarted, the A,nalysis phase identifies *'Tl* and *T3* as transactions

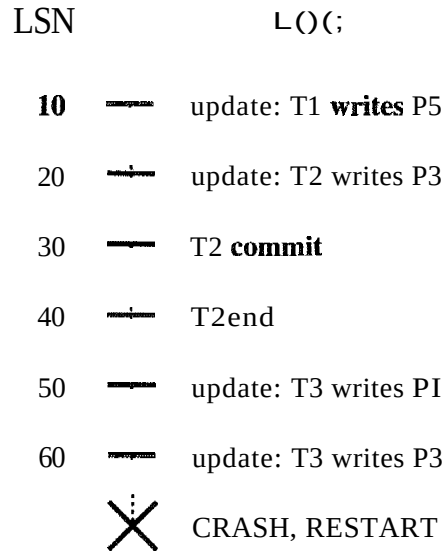


Figure 18.1 Execution History with a Crash

active at the time of the crash and therefore to be undone; *T2* as a corrnuitted transaction, and all its actions therefore to be written to disk; and PI, *P3,* and *P5* as potentially dirty pages. All the updates (including those of TI and *T3)* are reapplied in the order shown during the Redo phase. Finally, the actions of TI and *T3* are undone in reverse order during the Undo phase; that is, *T3's* write of *P3* is undone, *T3*'s write of PI is undone, and then TI's write of *P5* is undone.

Three Inain principles lie behind the ARIES recovery algoritlun:

- Write-Ahead Logging: Any change to a database object is first recorded in the log; the record in the log lllUst be written to stable storage before the change to the database object is written to disk.
- Repeating History During Redo: On restart following a crash, ARIES retraces all actions of the DBMS before the crash and brings the systern back to the exact state that it was in at the time of the crash. Then, it undoes the actions of transactions still active at the tirne of the crash (effectively aborting theln).
- Logging Changes During Undo: Changes lnada to the database while undoing a transaction are logged to ensure such an action is not repeated in the event of repeated (failures causing) restarts.

The second point distinguishes ARIES frorn other recovery algorithrns and is the basis for rnuch of its sirnplicity and flexibility. In particular, ABIES can support concurrency control protocols that involve locks of finer granularity than a page (e.g., record-level locks). The seCOJlc! and third points are also

Crash Recovery: IBM DB2, Inforrnix, Microsoft SQL Server, Oracle 8, and Sybase l\SE all use a WAL seherue for recovery. IBIvI DB2 uses ARIES, and the others use seherues that are actually quite sinlilar to ARIES (e.g., all changes are re-applied, not just the changes made by transactions that are 'winners') although there are several variations.

important in dealing with operations where redoing and undoing the operation are not exact inverses of each other. We discuss the interaction between concurrency control and crash recovery in Section 18.8, where we also discuss other approaches to recovery briefly.

18.2 THELOG

The log, SOlnetirnes called the trail or journal, is a history of actions executed by the DBMS. Physically, the log is a file of records stored in stable storage, which is assumed to survive crashes; this durability can be achieved by maintaining two or more copies of the log on different disks (perhaps in different locations), so that the chance of all copies of the log being sinlultaneously lost is negligibly small.

The most recent portion of the log, called the log tail, is kept in nlain Inemory and is periodically forced to stable storage. This way, log records and data records are written to disk at the same granularity (pages or sets of pages).

Every log record is given a unique *id* called the log sequence number (LSN). As with any record id, we can fetch a log record with one disk access given the LSN. Further, LSNs should be assigned in ruonotonically increasing order; this property is required for the ARIES recovery algorithrn. If the log is a sequential file, in principle growing indefinitely, the LSN can sirllply be the address of the first byte of the log record.[1]

For recovery purposes, every page in the database contains the LSN of the rnost recent log record that describes a change to this page. This LSN is called the pageLSN.

A log record is\vritten for each of the following actions:

[1] In practice, various techniques are used to identify portions of the log that are 'too old' to be needed again to bound the amount of stable storage used for the log. Given such a bound, the log may be implemented as a 'circular' file, in which case the LSN may be the log record id plus a *wrap-count*.

- Updating a Page: After rTlodifying the page, an *update* type record (described later in this section) is appended to the log tail. Tlhe pageLSN of the page is then set to the LSN of the update log record. (The page Blust be pinned in the buffer pool while these actions are carried out.)

- Conl1nit: When a transaction decides to conunit, it force-writes a *commit* type log record containing the transaction id. That is, the log record is appended to the log, and the log tail is written to stable storage, up to and including the cOllunit record.[2] The transaction is considered to have cOIlnnitted at the instant that its cOlnmit log record is written to stable storage. (Solne additional steps rnust be taken, e.g., reilloving the transaction's entry in the transaction table; these follow the writing of the cOlInnit log record.)

- Abort: When a transaction is aborted, an *abort* type log record containing the transaction id is appended to the log, and Undo is initiated for this transaction (Section 18.6.3).

- End: As noted above, when a transaction is aborted or comrnitted, some additional actions rnust be taken beyond writing the abort or COlllIllit log record. After all these additional steps are c()lnpleted, an *end* type log record containing the transaction id is appended to the log.

- Undoing an update: When a transaction is rolled back (because the transaction is aborted, or during recovery frorn a crash), its updates are undone. When the action described by an update log record is undone, a *cornpensation log record*, or CLR, is written.

Every log record has certain fields: prevLSN, transID, and type. The set of all log records for a given transaction is rnaintained as a linked list going back in tirne, using the prevLSN field; this list HUlst be updated whenever a log record is added. The transII) field is the id of the transaction generating the log record, and the type field obviously indicates the type of the log record.

Additional fields depend on the type of the log record. We already rnentioned the additional contents of the various log record types, with the exception of the update and compensation log record types, which we describe next.

Update Log Records

The fields in an update log record are illustrated in Figure 18.2. frhe pageID field is the page iel of the Inodified page; the length in bytes and the offset of the

[2]Note that this step requires the buffer manager to be able to selectively *force* pages to stable storage.

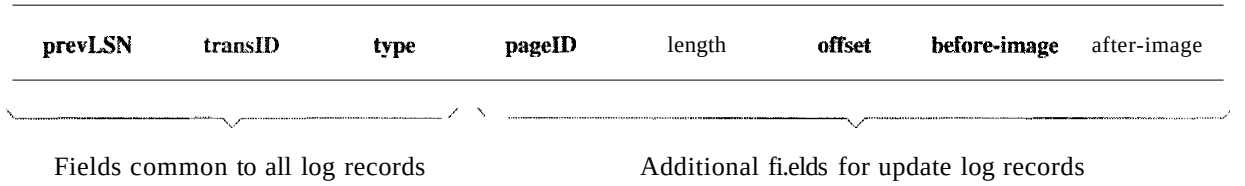


Figure 18.2 Contents of an Update Log Record

change are also included. The before-image is the value of the changed bytes before the change; the after-image is the value after the change. An update log record that contains both before- and after-images can be used to redo the change and undo it. In certain contexts, which we do not discuss further, we can recognize that the change will never be undone (or, perhaps, redone). A redo-only update log record contains just the after-iluage; similarly an undo-only update record contains just the before-iluage.

Compensation Log Records

A compensation log record (CLR) is written just before the change recorded in an update log record U is undone. (Such an undo can happen during nor-rnal system execution when a transaction is aborted or during recovery froIn a crash.) A cOlnpensation log record C describes the action taken to undo the actions recorded in the corresponding update log record and is appended to the log tail just like any other log record. 'fhe cornpensation log record C also contains a field called undoNextLSN, which is the LSN of the next log record that is to be undone for the transaction that wrote 11pdate record U; this field in C is set to the value of prevLSN in *[J.*

As an exarllple, consider the fourth update log reeord shown in Figure 18.3. If this update is undone, a CLR would be written, and the inforrnation in it would include the transII), pageID, length, offset, and before-image fields froln the update record. Notice that the CLR records the (undo) action of changing the affected bytes back to the before-irnage value; thus, this value and the location of the affected bytes constitute the redo infonnation for the action described by the CLR. The undoNextLSN field is set to the LSN of the first log record in Figure 18.:3.

lJnlike an update log record, a CLR describes an action that \vill never be *undone*, that is, we never undo an undo action. '1'he reason is sirnple: An update log record describes a change lnade by a transaction during nonnal execution and the transaction rnay subsequently be aborted, whereas a CLR describes an a,ctiol1 taken to rollback a transaction for which the decision to abort has already been rnade. Therefore, the transaction *must* be rolled back, and the

undo action described by the CLR is definitely required. This observation is very useful because it bounds the a1nount of space needed for the log during restart froin a crash: The nUlnber of CLRs that ca,n be written during LJndo is no 1nore than the nurnber of update log records for active transactions at the tirne of the crash.

A CLR Il1ay be written to stable stora,ge (follo\ving WAL, of course) but the undo action it describes rIlay not yet been written to disk when the systenl crashes again. In this case, the undo action described in the CLR is reapplied during the Redo phase, just like the action described in update log records.

For these reasons, a CLR contains the infonnation needed to reapply, 01' redo, the change described but not to reverse it.

18.3 OTHER RECOVERY-RELATED STRU'CTURES

In addition to the log, the following two tables contain important recovery-related infornlation:

- Transaction Table: This table contains one entry for each active transaction. 'The entry contains (arnong other things) the transaction id, the status, and a field called **lastLSN,** which is the LSN of the rnost recent log record for this transaction. The status of a transaction can be that it is in progress, corunlitted, or aborted. (In the latter two cases, the transaction will be rernoved fro1l1 the table once certain 'clean up' steps are c(nupleted.)
- Dirty page table: This table contains one entry for each dirty page in the buffer pool, that is, each page with changes not yet reflected on disk. The entry contains a field recLSN, which is the LSN of the first log record that caused the page to becorne dirty. Note that this LSN identifies the earliest log record that lnight have to be redone for this page during restart fronl a crash.

I)uring norrnal operation, these are rnainta..ined by the transaction rnanager and the buffer rnanager, respectively, and during restart after a crash, these tables are reconstructed in the Analysis phase of restart.

Consider the follc)\ving silupic exarnple. Transaction TIOOO changes the value of bytes 21 to 23 011 page P500 frorn 'ABC' to 'DEF', transaction 'T2000 changes 'HIJ' to 'KLM' on page P600, transaction $T2000$ changes bytes 20 through 22 fronl 'GDE' to 'QRS' on page $P500$, then transaction T1000 changes 'TUV' to 'WXY' on pageP505. The dirty page table, the transaction table,[3] and

[3]The status field is not shown in the figure for space reasons; all transactions are in progress.

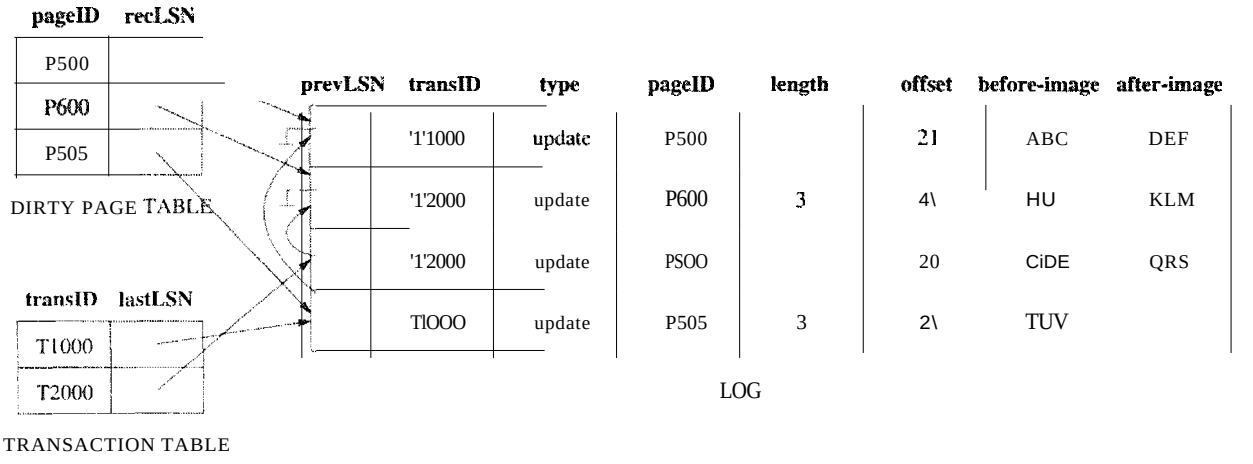


Figure 18.3 Instance of Log and Ttansaction Table

the log at this instant are shown in Figure 18.3. ()bserve that the log is shown growing froni top to bottorn; older records are at the top. Although the records for each transaction are linked using the prevLSN field, the log as a whole also has a sequential order that is iInportant---for exarnple, T2000's change to page P500 follows TlOOO's change to page *P500,* and in the event of a crash, these changes nUlst be redone in the sanle order.

18.4 THE WRITE-AHEAD LOG **PROTOCOL**

Before writing a page to disk, every update log record that describes a change to this page rnust be forced to stable storage. This is accornplished by forcing all log records up to and including the one with LSN equal to the pageLSN to stable storage before writing the page to disk.

The irnportance of the WAL protocol carulot be overerl1phasized- --\VAL is the fundarnentaJ rule that ensures that a record of every change to the database is available while atternpting to recover froni a crash. If a transaction rnade (l. change and committed, the no-force approach Incans that some of these changes rnay not have been written to disk at the tirne of a sulJsequent crash. Without a record of these changes, there would be no way to ensure that the changes of a corn11.itted transaction survive crashes. Note that the definition of a *committed transaction* is effectively 'a transa,ction all of whose log records, including a conunit record, have l)een written to stable storage'.

When a transaction is cornrnitted, the log tail is forced to stable storage, even if a no-force approach is being used. It is worth contrasting this operation with the a,ctions taken under a force approach: If a force approach is used, all the pages rIlodified by the transaction, rather than a portion of the log that includes all its records, IHIIS!, be forced to disk when the transaction conllIlits. The set of

all changed pages is typically 11luch larger than the log tail because the size of an update log record is close to (twice) the size of the changed bytes, ·which is likely to be Inuch s1na11er than the page size. Further, the log is 1naintained as a sequential file, and all writes to the log are sequential writes. Consequently, the cost of forcing the log tail is luuch sIllaller than the cost of \vriting all changed pages to disk.

18.5 CHECKPOINTIN(;

A checkpoint is like a snapshot of the DBMS state, and by taking checkpoints periodically, as we will see, the DBl\1S can reduce the alnount of work to be done during restart in the event of a subsequent crash.

Checkpointing in ARIES has three steps. First, a begin_checkpoint record is written to indicate when the checkpoint starts. Second, an end_checkpoint record is constructed, including in it the current contents of the transaction table and the dirty page table, and appended to the log. The third step is carried out after the end_checkpoint record is written to stable storage: A special master record containing the LSN of the *begirLcheckpoint* log record is written to a known place on stable storage. While the end_checkpoint record is being constructed, the DBMS continues executing transactions and writing other log records; the only guarantee we have is that the transaction table and dirty page table are accurate *as of the time of the begin_checkpoint record.*

This kind of checkpoint, called a fuzzy checkpoint, is inexpensive because it does not require quiescing the SystCIll or writing out pages in the buffer pool (unlike some other forlns of checkpointing). On the other hand, the effectiveness of this checkpointing technique is lirnited by the earliest recLSN of pages in the d.irty pages table, because during restart we Inust redo changes starting froin the log record \vhose LSN is equal to this recI.lSN. l-Iaving a background process that periodically writes dirty pages to disk helps to lirnit this probleln.

When the SystCIll cornes back up after a crash, the restart process begins by locating the rnost recent checkpoint record. For uniforlnity, the systeIll always begins no1'n1al execution by takirlg a checkpoint, in which the transaction table and dirty page table are both Clnpty.

18.6 RECOVERING FROM A SYSTEM CRASH

\Vhen the systenl is restarted after a crash, the recovery manager proceeds in three phases, as shown in Figure 18.4.

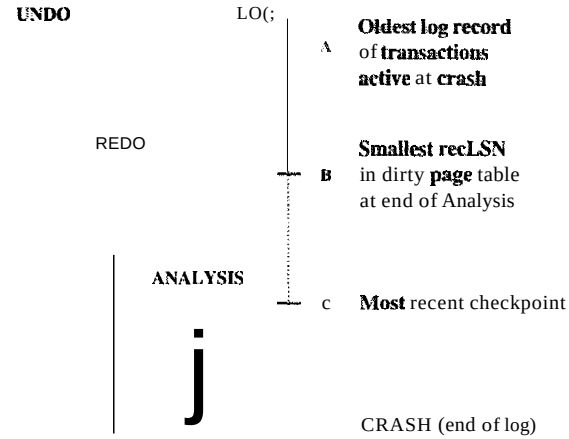


Figure 18.4 Three Phases of Restart in ARIES

The Analysis phase begins by examInIng the rnost recent begin_checkpoint record, whose LSN is denoted C in Figure 18.4, and proceeds forward in the log until the last log record. '1'he Redo phase follows Analysis and redoes all changes to any page that Illight have been dirty at the tirl1e of the crash; this set of pages and the starting point for Redo (the srnallest recLSN of any dirty page) are deterrnined during Analysis. 'The Undo phase follows Redo and undoes the changes of all transactions active at the tirne of the crash; again, this set of transactions is identified during the Analysis phase. Note that Redo reapplies changes in the order in which they were originally carried out; Undo reverses changes in the opposite order, reversing the lllost recent change first.

Observe that the relative order of the three points *A*, *B*, and C in the log rnay differ frolIl that shown in Figure 18.4. The three phases of restart are described in rnore detail in the following sections.

18.6.1 Analysis Phase

The Analysis phase perfonns three tasks:

1. It detennines the point in the log at which to start the Redo pass.
2. It deterrnines (a conservative superset of the) pages in the buffer pool that were clirty at the tirne of the crash.
3. It identifies transactions that were active at the tirne of the crash and rnust be undone.

Analysis begins by exEtrnining the rnost recent begirLcheckpoint log record and initializing the dirty page table and transaction table to the copies of those structures in the next end_checkpoint record. Thus, these tables are initialized to the set of dirty pages and active transactions at the tilne of the checkpoint.

(If additional log records are between the begiILcheckpoint and encLcheckpoint records, the tables HIUst be adjusted to reflect the inforluation in these records, but \ve cnnit the details of this step. See Exercise 18.9.) A.naJysis then scans the log in the for\vard direction until it reaches the end of the log:

- If an end log record for a transaction *T* is encountered,T is reIlloved fronl the transaction table because it is no longer active.
- If a log record other than an end record for a transaction *T* is encountered, an entry for *T* is added to the transaction table if it is not already there. Further, the entry for *T* is rnodified:
  1. The lastLSN field is set to the LSN of this log record.
  2. If the log record is a cOllnnit record, the status is set to C, otherwise it is set to U (indicating that it is to be undone).
- If a redoable log record affecting page *P* is encountered, and *P* is not in the dirty page table, an entry is inserted into this table with page id *P* and recLSN equal to the LSN of this redoable log record. This LSN identifies the oldest change affecting page *P* that may not have been written to disk.

At the end of the Analysis phase, the transaction table contains an accurate list of all transactions that were active at the tilue of the crash—this is the set of transactions with status U. The dirty page table includes all pages that were dirty at the tirne of the crash but rnay also contain SOlne pages that were written to disk. If an *end_write* log record were written at the cornpletion of ea,ch write operation, the dirty page table constructed during Analysis could be lnade rnore accurate, but in AHJES, the additional cost of writing eneLwrite log records is not considered to be worth the gain.

As an example, consider the execution illustrated in Figure 18.3. Let us extend this execution by assurning that *T*'2000 COlIllnits, then *TIOnO* rnodifies another page, say, .P700, and appends an update record to the log tail, and then the systern crashes (before this update log record is written to stable storage).

The dirty page table and the transaction table, held in rnernory, are lost in the crash. The rnost recent checkpoint was taken at the beginning of the execution, \vith an ernpty tran.saction table and dirty page table; it is not shown in Figure 18.3. After examining this log record, \vhich we assurne is just before the first log record shown in the figure, Analysis initializes the two tables to l>e ernpty. Scanning forward in the log, T'1000 is added to the transaction table; in additiol1,P500 is added to the dirty page ta,blc\vith recLSN equal to the LSN of the first sho\vn log record. Sirnilarly, T2C)OO is added to the transaction table andPGOO is added to the dirty page table. There is no change based on the third log record, and the fourth record results in the addition of *P505* to

the dirty page table. The eOllnnit record forT2000 (not in the figure) is no\v encountered, and T2000 is relIloved fro111 the transaction table.

The Analysis phase is now eornplete, and it is recognized that the only active transaction at the time of the crash is *TIOOO,* \vith lastLSN equal to the LSN of the fourth record in Figure 18.3. rrhe dirty page table reconstructed in the Analysis phase is identical to that shown in the figure. The update log record for the change to *P700* is lost in the crash and not seen during the Analysis pass. Thanks to the WAL protocol, however, all is well·····--thecorresponding change to page *P700* cannot have been written to disk either!

SaIne of the updates rnay have been written to disk; for concreteness, let us assume that the change to *P600* (and only this update) was written to disk before the crash. ThereforeP600 is not dirty, yet it is included in the dirty page table. rIhe pageLSN on page *P600,* however, reflects the write because it is now equal to the LSN of the second update log record shown in Figure 18.3.

18.6.2 Redo Phase

During the Redo phase, ARIES reapplies the updates of *all* transactions, COIll-rnitted or otherwise. Further, if a transaction was aborted before the crash and its updates were undone, as indicated by CLRs, the actions described in the CLRs are also reapplied. This repeating history paradigm distinguishes ARIES from other proposed vVAL-based recovery algoritlnIls and causes the database to be brought to the sarne state it was in at the time of the crash.

rrhe R,edo phase begins with the log record that has the srnallest recLSN of all pages in the dirty page table constructed by the Analysis pass because this log record identifies the oldest update that rnay not have been written to disk prior to the crash. Starting frorn this log record, R,edo scans forward until the end of the log. For each redoable log record (update or CLR) encountered, Redo checks whether the logged action HUlst be redone. The action rnust be redone unless one of the follo\ving conditions holds:

- The affected page is not in the dirty page table.
- rrhe affected page is in the dirty page table, but the recLSN for the entry is *greater than* the LSN of the log record being checked.
- 1'he pageLSN (stored on the page, which rnust be retrieved to check this condition) is *greater than or equal* to the LSN of the log record being checked.

The first condition obviously 1118a11S that all changes to this page have been written to disk. Because the recLSN is the first update to this page that lnay

not have been written to disk, the second condition rneans that the update being checked was indeed propagated to disk. The third condition, which is checked last because it requires us to retrieve the page, also ensures that the update being checked was ·written to disk, because either this update or a later update to the page was written. (Recall our assumption that a write to a page is atomic; this assurnption is irnportant here!)

If the logged action lllust be redone:

1. The logged action is reapplied.
2. The pageLSN on the page is set to the LSN of the redone log record. No additional log record is written at this time.

Let us continue with the exarnple discussed in Section 18.6.1. FrorIl the dirty page table, the smallest recLSN is seen to be the LSN of the first log record shown in Figure 18.3. Clearly, the changes recorded by earlier log records (there happen to be none in this example) have been written to disk. Now, Redo fetches the affected page, *P500,* and compares the LSN of this log record with the pageLSN on the page and, because we assurned that this page was not written to disk before the crash, finds that the pageLSN is less. The update is therefore reapplied; bytes 21 through 23 are changed to 'DEF', and the pageLSN is set to the LSN of this update log record.

Redo then exarnines the second log record. Again, the affected page, P600, is fetched and the pageLSN is cornpared to the LSN of the update log record. In this case, because we assurned thatP600 was written to disk before the crash, they are equal, and the update does not have to be redone.

The rernaining log records are processed sirnilarly, bringing the systern back to the exact state it was in at the tirue of the crash. Note that the first two conditions indicating that a redo is unnecessary never hold in this exaruple. Intuitively, they corne into play when the dirty page table contains a very old recLSN, going back to before the rJlost recent checkpoint. In this case, as Redo scans forwa.rd frorn the log record with this LSN, it encounters log records for pages that were written to disk prior to the checkpoint and therefore not in the dirty page table in the checkpoint. Sorne of these pages Inay be dirtied again after the checkpoint; nonetheless, the updates to these pages prior to the checkpoint need not be redone. Although the third condition alone is sufficient to recognize that these updates need not be redone, it requires us to fetch the affected page. The first t\VO conditions allow us to recognize this situation \vithout fetching the page. (The reader is encouraged to construct exaulples that illustrate the use of each of these conditions; see Exercise 18.8.)

At the end of the Redo phase, end type records are written for all transactions with status C, which are rCllloved '£1'01n the transaction table.

18.6.3 Undo Phase

The Undo phase, unlike the other two phases, scans backward fronl the end of the log. The goal of this phase is to undo the actions of all transactions active at the tilne of the crash, that is, to effectively abort the1n. This set of transactions is identified in the transaction table constructed by the AIlalysis phase.

The Undo Algorithm

Undo begins with the transaction table constructed by the .Analysis phase, which identifies all transactions active at the time of the crash, and includes the LSN of the 1110st recent log record (the lastLSN field) for each such transaction. Such transactions are called loser transactions. All actions of losers IllUst be undone, and further, these actions rnust be undone in the reverse of the order in which they appear in the log.

Consider the set of lastLSN values for all loser transactions. Let us call this set ToUndo. Undo repeatedly chooses the largest (Le., rnost recent) LSN value in this set and processes it, until rro1Jndo is ernpty. To process a log record:

1. If it is a CLR and the undoNextLSN value is not *null*, the undoNextLSN value is added to the set ToUndo; if the undoNextLSN is *null*, an end record is written for the transaction because it is cornpletely undone, and the CLR, is discarded.
2. If it is an. update record, a CLR, is written and the corresponding a,ction is undone, as described in Section 18.2, and the prevLSN value in the update log record is added to the set ToUndo.

When the set rroUndo is empty, the lJndo phase is cornplete. I{estart is no\v cornplete, and the systenl can proceed with nonnal operations.

Let us continue with the scenario discussed in Sections 18.6.1 and 18.6.2. The only active transaction at the tiTne of the crash was detennined to be TI000. :F'rom the transaction table, we get the LSN of its Inost recent log record, which is the fourth update log record in Figure 18.3. '1'he update is undone, and a CLR is \vritten\vith undoNextLSN equal to the LSN of the first log record in the figure. The next record to be undone for transaction $T1000$ is the first log record in the figure. After this is undone, a CLR a,nel an end log record for $T1000$ are written, and the IJndo phase is cornplete.

In this exarnple, undoing the action recorded in the first log record causes the action of the third log record, \vhich is due to a conunitted traJlsaetioIl, to be overwritten and thereby lost! rrhis situation arises because $T2000$ overwrote a data itcrIl \vritten by TIOOO while $T1000$ was still active; if Strict 2PLwere followed, $T2000$ would not have been allowed to overwrite this data iterH.

Aborting a Transaction

Aborting a transaction is just a special case of the Undo phase of Restart in which a single transaction, rather than a set of transactions, is undone. The exarnple in Figure 18.5, discussed next, illustrates this point.

Crashes during Restart

It is important to understand how the lTndo algorithrn presented in Section 18.6.3 handles repeated systern crashes. Because the details of precisely how the action described in an update log record is undone are straightforward, we discuss Undo in the presence of systern crashes using an execution history, shown in Figure 18.5, that abstracts away unnecessary detail. This exarnple illustrates how aborting a transaction is a special case of Undo and how the use of CLRs ensures that the Undo action for an update log record is not applied twice.

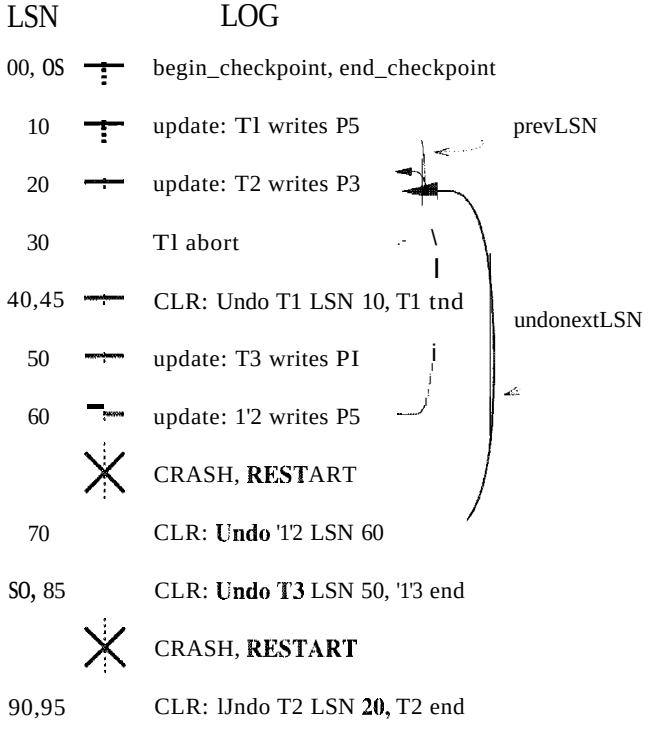


Figure 18.5 Example of Undo with Repeated Crashes

The log shows the order in which the DBMS executed various actions; note that the LSNs are in ascending order, and that each log record for a transaction has a prevLSN' field that points to the previous log record for that transaction. We have not shown *null* prevLSNs, that is, SOIne special value used in the prevLSN field of the first log record for a, transaction to indicate tha,t there is no previous log record. We also cOlnpacted the figure by occasionally displaying two log records (separated by a cOIlllna) on a single line.

Log record (with LSN) 30 indicates that *T1* aborts. All actions of this transaction should be undone in reverse order, and the only action of $T1$, described by the update log record 10, is indeed undone as indicated by CLR, 40.

After the first crash, Analysis identifies F)l (with recLSN 50), $P3$ (with recLSN 20), and *P5* (with recLSN 10) as dirty pages. Log record 45 shows that *T1* is a cornpleted transaction; hence, the transaction table identifies *T2* (with lastLSN 60) andT3 (with lastLSN 50) as active at the tirne of the crash. '1'he Redo phase begins with log record 10, which is the rninirnurn recLSN in the dirty page table, and reapplies all actions (for the update and CLR, records), as per the Redo algorithIl1 presented in Section 18.6.2.

The r1'olJndo set consists of LSNs 60, for $T2$, and 50, for $T3$. The lJndo phase now begins by processing the log record with LSN 60 because 60 is the largest LSN in the ToUndo set. The update is undone, and a CLR, (with LSN 70) is written to the log. This CLR has llndoNextLSN equal to 20, which is the prevLSN value in log record 60; 20 is the next action to be undone for 1 2. Now the largest rernaining LSN in the rrOUndo set is 50. The write corresponding to log record 50 is now undone, and a CLH, describing the change is 'written. rrhis CLR has LSN 80, and its undoNextLSN field is *null* because 50 is the only log record for transaction *T3*. Therefore $T3$ is cOITlpletely undone, and an end record is written. Log records 70, 80, and 85 are written to stable storage before the systern crashes a second tirHe; however, the changes described by these records ITlay not have been written, to disk..

When the systern is resta.rted after the S8C011(1 crash. Analysis deterrnines that the only active transactioIl at the time of the crash was *T2;* in addition, the dirty pa,ge table is identical to what it was during the previous restart. Log records 10 througll 85 are processed again during Redo. (If sorne of the changes made during the previous Redo were written to disk, the pageLSN's on the affected pages are used to detect this situation and avoid writing these pages again.) The lJndo phase considers the onlyLSN in the TolJndo set, 70, and processes it }))" adding tIle llndoNextLSN value (20) to the ToUndo set. Next, log record 20 is processed l)y undoing $T2$'s write of page $P3$, a.nd a CLR is written (LSN 90). Because 20 is the first of 7'2's log records and therefore, the last of its records

to be undone—the undoNextLSN field in this CLR is *null*, an end record is written for *T2*, al1d the TolJndo set is now empty.

Recovery is no\v complete, and norrnal execution can resurne with the ·writing of a checkpoint record.

This exarnple illustrated repeated crashes during the lJndo phase. For cornpleteness, let us consider what happens if the system crashes while R,estart is in the Analysis or Redo phase. If a crash occurs during the Analysis phase, all the work done in this phase is lost, and on restart the Analysis phase starts afresh with the sa11le inforrnation as before. If a crash occurs during the Redo phase, the only effect that survives the crash is that sorne of the changes rnade during Redo may have been written to disk prior to the crash. R,estart starts again with the Analysis phase and then the Redo phase, and sorne update log records that were redone the first tirne around will not be redone a second tirne because the pageLSN is now equal to the update record's LSN (although the pages have to be fetched again to detect this).

We can take checkpoints during Restart to rninirnize repeated work in the event of a crash, but we do not discuss this point.

18.7 MEDIA RECOVERY

Media recovery is based on periodically rnaking a copy of the database. Because copying a large database object such as a file can take a long tirHe, and the I)BMS rnust be allowed to continue with its operations in the Ineantirne, creating a copy is handled in a rnanner sirnilar to taking a fuzzy checkpoint.

When a database object such as a file or a page is corrupted, the copy of that object is brought up-to-date by using the log to identify and reapply the changes of cornnlitted transactions and undo the changes of uncoI1unitted transactions (as of the tirne of the rnedia recovery operation).

The begin_checkpoint LSN of the rnost recent cOlllplete checkpoint is recorded along with the copy of the database object to luinirnize the work in reapplying changes of committed transactions. Let us COlnpare the smallest recLSN of a dirty page in the corresponding encLcheckpoint record \vith the I;SN of the begirLcheckpoint record and call the slua.ller of these two LSNs *I*. We observe that the actions recorded in all log records with LSNs less than *I* Inust be reflected in the copy. Thus, 0111y log records with LSNs greater than *I* need be reapplied to the copy.

Finally, the updates of transactions that are incornplete at the tiIl1e of Inedia recovery or that were aborted after the fuzzy copy was corllpleted need to be undone to ensure that the page reflects only the actions of conunitted transactions. The set of such transactions can be identified as in the Analysis pass, and we ornit the details.

18.8 OTHER APPROACHES AND INTERACTION WITH CONCURRENCY CONTROL

Like ARIES, the Inost popular alternative recovery algoritlllns also rnaintain a log of database actions according to the WAL protocol. A InajaI' distinction between ARIES and these variants is that the Redo phase in ARIES *repeats history,* that is, redoes the actions of *all* transactions, not just the non-losers. Other algorithms redo only the non-losers, and the Redo phase follows the Undo phase, in which the actions of losers are rolled back.

Thanks to the repeating history paradigm and the use of CLRs, ARIES supports fine-granularity locks (record-level locks) and logging of logical operations rather than just byte-level rnodifications. For exalllple, consider a transaction *T* that inserts a data entry 15* into a B+ tree index. Between the time this insert is done and the time that *T* is eventually aborted, other transactions Inay also insert and delete entries frorn the tree. If record-level locks are set rather than page-level locks, the entry 15* I11ay be on a different physical page when *T* aborts fr0111 the one that *T* inserted it into. In this case, the undo operation for the insert of 15* lllUSt be recorded in logical tenns because the physical (byte-level) actions involved in undoing this operation are not the inverse of the physical actions involved in inserting the entry.

Logging logical operations yields considerably higher concurrency, although the use of fine-granularity locks can lead to increased locking activity (because rnore locks 1nust be set). Hence, there is a trade-off between different WAL-based recovery schclnes. We chose to cover ARIES because it has several attractive properties, in particular, its sirnplicity and its ability to support fine-granularity locks and logging of logical operations.

One of the earliest recovery algorithrns, llsed in the Syster11 R prototype at IBI\'1, takes a very different approach. 'There is no logging and, of course, no WAL protocol. Instead, the database is treated as a collection of pages and accessed thTough a page table, which maps page ids to disk addresses. When a transaction Inakes changes to a data page, it actually Inakes a copy of the page, called the shadow of the page, a,nel changes the shadow page. The transaction copies the appropriate part of the page table and chan,ges the entry for the changed page to point to the shadow, so that it can see the

changes; ho\vever, other transactions continue to see the original page table, and therefore the original page, until this transaction COll1lnits. Aborting a transaction is sirnple: .Just discard its shadow versions of the page table and the data pages. Cornrnitting a transaction involves rnaking its version of the page table public and discarding the original data pages that are superseded by shado\v pages.

This schelue suffers frorn a nUlnber of problerlls. First, data becornes highly fragrnented clue to the replacernent of pages by shadow versions, which rIlay be located far fr01n the original page. This phenornenon reduces data clustering and rnakes good garbage collection irnperative. Second, the sche1ne does not yield a sufficiently high degree of concurrency. rrhird, there is a substantial storage overhead due to the use of shadow pages. Fourth, the process aborting a transaction can itself run into deadlocks, and this situation rllust be specially handled because the sernantics of aborting an abort transaction gets rnurky.

For these reasons, even in Systern R, shadow paging was eventually superseded by \VAL-based recovery techniques.

18.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are the advantages of the ARIES recovery algoritluu? (Section 18.1)
- Describe the three steps in crash recovery in ARIES? What is the goal of the Analysis phase? The redo phase? The undo phase? (Section 18.1)
- What is the LSN of a log record? (Section 18.2)
- \Vhat are the different types of log records and when are they written? (Section 18.2)
- What inforrnation is rnaintained in the transaction table and the dirty page table? (Section 18.3)
- What is Write-Ahead Logging? \Vhat is forced to disk at the tirne a transaction COIlllnits? (Section 18.4)
- What is a fuzzy checkpoint? \Vhy is it useful? What is a master log record? (Section 18.5)
- In \vhich direction does the .A.nalysis phase of recovery scan the log? At \vhich point in the log does it begin and end the scan? (Section 18.6.1)
- Descril)c \vhat infonnation is gathered in the Analysis phase and ho\v. (Section 18.6.1)

- In \vhich direction does the Redo phase of recovery process the log? At which point in the log does it begin and end? (Section 18.6.2)
- What is a redoable log record? Under what conditions is the logged action redone? \Vhat steps are carried out when a logged action is redone? (Section 18.6.2)
- In which direction does the Undo phase of recovery process the log? At which point in the log does it begin and end? (Section 18.6.3)
- What are loser transactions? How are they processed in the Undo phase and in what order? (Section 18.6.3)
- Explain what happens if there are crashes during the Undo phase of recovery. What is the role of CLRs? What if there are crashes during the Analysis and Redo phases? (Section 18.6.3)
- How does a DBMS recover from 111edia failure without reading the complete log? (Section 18.7)
- Record-level logging increases concurrency. What are the potential problems, and how does ARIES address them? (Section 18.8)
- What is shadow paging? (Section 18.8)

EXERCISES

Exercise 18.1 Briefly answer the following questions:

1. How does the recovery rnanager ensure atornicity of transactions? How does it ensure durability?
2. What is the difference between stable storage and disk?
3. What is the difference between a systenl crash and a media failure?
4. Explain the WAL protocol.
5. Describe the steal and no-force policies.

Exercise 18.2 Briefly answer the follO\ving questions:

1. What are the properties required of LSNs?
2. What are the fields in an update log record? Explain the use of each field.
3. VVhat are redoal)le log records?
4. What are the differences between update log records and CLRs?

Exercise 18.3 Briefly answer the following questions:

1. What are the roles of the Analysis, Redo, and Undo phases in ARIES?
2. Consider the execution shown in Figure 18.6.

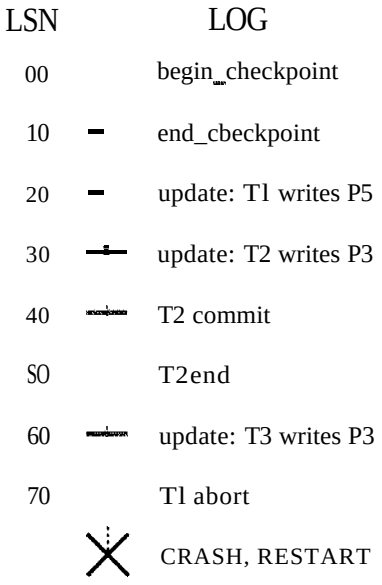


Figure 18.6 Execution with a Crash

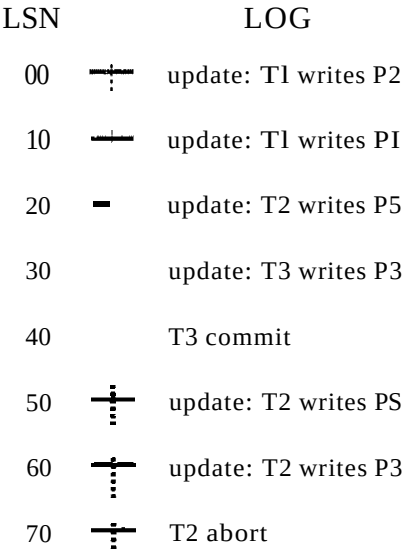


Figure 18.7 Aborting a Transaction

(a) What is done during Analysis? (Be precise about the points at which Analysis begins and ends and describe the contents of any tables constructed in this phase.)

(b) What is done during Redo? (Be precise about the points at which Redo begins and ends.)

(c) What is done during Undo? (Be precise about the points at which Undo begins and ends.)

Exercise 18.4 Consider the execution shown in Figllre 18.7.

1. Extend the figure to shuw prevLSN and llndonextLSN values.
2. Describe the actions taken to rollback transaction $T2$.

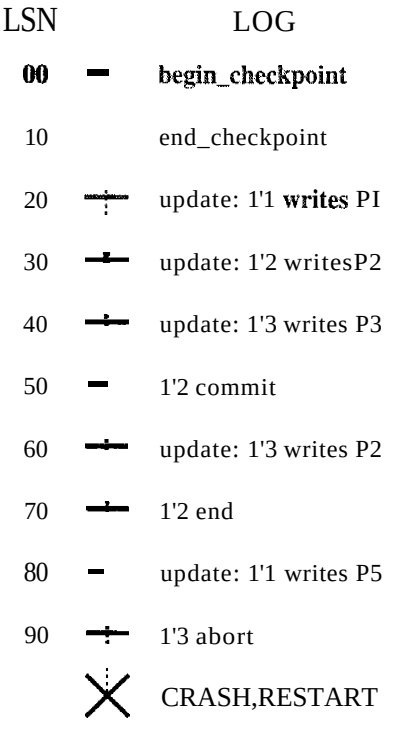


Figure 18.8 Execution with Multiple Crashes

3. Show the log after *T2* is rolled back, including all prevLSN and undonextLSN values in log records.

Exercise 18.5 Consider the execution shown in Figure 18.8. In addition, the systerll crashes during recovery after writing two log records to stable storage and again after writing another two log records.

1. What is the value of the LSN stored in the master log record?
2. What is done during Analysis?
3. What is done during Redo?
4. \What is done during Undo?
5. Show the log when recovery is complete, including all non-null prevLSN and unclonextLSN values in log records.

Exercise 18.6 Briefly answer the following questions:

1. How is checkpointing done in ARIES?
2. Checkpointing can also be done as follows: Quiesce the systerll so that only checkpointing activity can be in progress, write out copies of all dirty pages, and include the dirty page table and trallsaction table in the checkpoint record. What are the pros and cons of this approach versus the checkpointiug a,pproach of ARIES?
3. What happens if a second begiILcheckpoint record is encountered during the Analysis phase?
4. C;an a second en(Lcheckpoint record be encountered during the AnaJysis phase?
5. Why is the use of CLRs irnportant for the use of undo actions that are not the physical inverse of the original update?

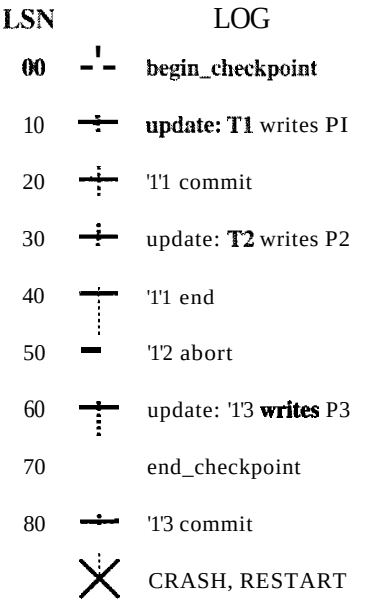


Figure 18.9 Log Records between Checkpoint Records

6. Give an example that illustrates how the paradigm of repeating history and the use of CLRs allow ARIES to support locks of finer granularity than a page.

Exercise 18.7 Briefly answer the following questions:

1. If the system fails repeatedly during recovery, what is the rrlaximum nunlber of log records that can be written (as a function of the number of update and other log records written before the crash) before restart cOInpletes successfully?
2. What is the oldest log record we need to retain?
3. If a bounded amount of stable storage is used for the log, how can we always ensure enough stable storage to hold all log records written during restart?

Exercise 18.8 Consider the three conditions under which a redo is unnecessary (Section 20.2.2).

1. \Vhy is it cheaper to test the first two conditions?
2. Describe an execution that illustrates the use of the first condition.
3. Describe an execution that illustrates the use of the second condition.

Exercise 18.9 The description in Section 18.6.1 of the Analysis phase rnade the sirnplifying assulTlptioll that no log records appeared between the begill-checkpoint and end_checkpoint records for the Inost recent cOlnplete checkpoint. The following questions explore how such records should be handled.

1. Explain why log records could be written between the begiIl-checkpoint and eneLcheckpoint records.
2. Describe how the Analysis phase could be Inodified to handle such records.
3. Consider the execution sho\vn in Figure 18.9. Show the contents of the encLcheckpoint record.
4. Illustrate your rnodified Analysis phase on the execution shown in Figure 18.9.

Exercise 18.10 Answer the following questions briefly:

1. Explain how 1nedin recovery is handled in ARIES.
2. What are the pros and cons of using fuzzy durnps for media recovery?
3. What are the sirYlilarities and differences between checkpoints and fuzzy chunps?
4. Contrast ARIES with other WAL-based recovery schernes.
5. Contrast AHlES with shadow-page-based recovery.

BIBLIOGRAPHIC NOTES

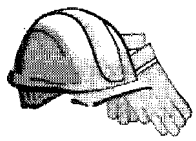
Our discussion of the ARIES recovery algorithm is based on [544]. [282] is a survey article that contains a very readable, short description of ARIES. [541, 545] also discuss ARIES. Fine,·granularity locking increases concurrency but at the cost of 11101'e locking activity; [542] suggests a technique based on LSNs for alleviating this problerYl. [458] presents a for111al verification of ARIES.

[355] is an excellent survey that provides a broader treatrnent of recovery algoritlulls than our coverage, in which we chose to concentrate on one particular algorithrn. [17] considers perfor-rnance of concurrency control and recovery algorithrIls, taking into account their interactions. The irnpact of recovery on concurrency control is also discussed in [769]. [625] contains a perforrnance analysis of various recovery techniques. [236] cornpares recovery techniques for main rnerllory database systeulS, which are optirnized for the case that 11l0st of the active data set fits in rnain H1ernory.

[478] presents a description of a recovery algorithm based on write-ahead logging in which 'loser' transactions are first undone and then (only) transactions that corl1nlitted before the crash are redone. Shadow paging is described in [493, 337]. A scherne that uses a cOlnbination of shadow paging and in-place updating is described in [624].

PART VI

DATABASE DESIGN AND TUNING



19

SCHEMA REFINEMENT AND NORMAL FORMS

- What problems are caused by redundantly storing information?
- What are functional dependencies?
- What are nornlal forms and what is their purpose?
- What are the benefits of BCNF and 3NF?
- What are the considerations in decolllposing relations into appropriate normal forms?
- Where does normalization fit in the process of database design?
- Are luore general dependencies useful in database design?
- Key concepts: redundancy, insert, delete, and update anomalies; functional dependency, Armstrong's Axioms; dependency closure, attribute closure; normal fonns, BCNF, 3NF; decOlnpositions, lossless-join, dependency-preservation; multivalued dependencies, join dependencies, inclusion dependencies, 4NF, 5NF

It is a nlelancholy truth that even great IneIl have their poor relations.

Charles Dickens

Conceptual database design gives us a set of relation schemas and integrity constraints (ICs) that can be regarded as a good starting point for the final database design. T'his initial design IHust be refined by taking the lCg into account rnore fully than is possiblc\vith just the ER rnodel constructs aIId also by considering perforrnance criteria and typical workloads. In this chapter, we c!iscllss how lCs can be used to refine the conceptual schema produced by

translating an ER 1Hodel design into a collection of relations. \Vorkload and perforrnance considerations are discussed in Chapter 20.

We concentrate on an irnportant class of constraints called *functional dependencies.* Other kinds of les, for exarnple, *multivalued dependencies* and *join dependencies*, also provide useful inforrnation. They can sOIuetilnes reveal redundancies that cannot be detected using functional dependencies alone. We discuss these other constraints briefly.

This chapter is organized as follows. Section 19.1 is an overview of the schenla refineInent approach discussed in this chapter. We introduce functional dependencies in Section 19.2. In Section 19.3, we show how to reason with functional dependency information to infer additional dependencies from a given set of dependencies. We introduce norlnal forIns for relations in Section 19.4; the normal form satisfied by a relation is a measure of the redundancy in the relation. A relation with redundancy can be refined by *decomposing it,* or replacing it with smaller relations that contain the saIne information but without redundancy. We discuss deco1npositions and desirable properties of decompositions in Section 19.5, and we show how relations can be decomposed into smaller relations in desirable normal forms in Section 19.6.

In Section 19.7, we present several examples that illustrate how relational schemas obtained by translating an ER model design can nonetheless suffer froln redundancy, and we discuss how to refine such schemas to eliminate the problems. In Section 19.8, we describe other kinds of dependencies for database design. We conclude with a discussion of nornlalization for our case study, the Internet shop, in Section 19.9.

19.1 INTRODUCTION TO SCHEMA REFINEMENT

We now present an overview of the probleIns that schenla refinement is intended to address and a refinernent approach based on decolnpositions. Iledundant storage of inforrnation is the root cause of these problerns. Although decoInposition can elirninate redundancy, it can lead to problclns of its own and should be used with caution.

19.1.1 Problems Caused by Redundancy

Storing the same inforrnation redundantly, that is, in l110re than one place \vithin a database, can lead to several problcll1S:

- Redundant Storage: SOU1C iuforInation is stored repeatedly.

- Update Anomalies: If one copy of sueh repeated data is updated, an inconsistency is created unless all copies are sirnilarly updated.
- Insertion Anomalies: It may not be possible to store certain inforlnation unless sorne other, unrelated, inforIIlatioIl is stored as well.
- Deletion Anomalies: It rnay not be possible to delete certain inforrnation without losing SOHle other, unrelated, infofrnation as well.

Consider a relation obtained by translating a variant of the IIourly_Emps entity set frorn Chapter 2:

Hourly_Emps(*ssn*, *name*, *lot*, *rating*, *hourly_wages*, *hours_worked*)

In this chapter, we ornit attribute type inforrnation for brevity, since our focus is on the grouping of attributes into relations. We often abbreviate an attribute narne to a single letter and refer to a relation schema by a string of letters, one per attribute. For exarllple, we refer to the Hourly_Ernps scherna as *SNLRWH* (*W* denotes the *hourly_wages* attribute).

1'he key for Hourly_Emps is *ssn.* In addition, suppose that the *hourly_wages* attribute is deterrnined by the *rating* attribute. That is, for a given *rating* value, there is only one perrllissible *hourly_wages* value. This IC is an exanlple of a *functional dependency.* It leads to possible redundancy in the relation Hourly_Ernps, as illustrated in Figure 19.1.

	I *narne*	*lot*	*rating*	*hourly_wages*	*hours_worked*
123-22-3666	Attishoo	48	8	10	40
231-31-5368	Sruiley	22	8	10	30
131-24-3650	Srllethurst	35	5	7	30
434-26-3751	Guldu	35	5	7	32
612-67-4134	Madayan	35	8	10	40

Figure 19.1 An Instance of the Hourly_Emps Relation

If the same value appears in the *rating* colurnn of two tuples, the IC tells us that the sarne value HUlst appear in the *hourly_wages* colurnn as well. This redundancy has the sarne negative consequences as before:

- *Redundant Storage:* rrhe rating value 8 corresponds to the hourly wage 10, and this association is repeated three tirnes.
- *[Tpdate Anomalies:* The *hourly_wages* in the first tuple could be updated without rnaking a sirnilar change in the second tuple.

- *Insertion Anomalies:* We cannot insert a tuple for an crnployee unless \ve know the hourly wage for the ernployee's rating value.
- *Delet'ion Anomalies:* If we delete all tuples with a given rating value (e.g., we delete the tuples for Snlcthurst and Guldu) we lose the association between that *rating* value and its *hourly_wage* value.

Ideally, we want schemas that do not pennit redundancy, but at the very least we want to be able to identify schernas that do allow redundancy. Even if we choose to accept a scherna with sorne of these drawbacks, perhaps owing to perforlnance considerations, we want to rnake an infonned decision.

Null Values

It is worth considering whether the use of *null* values can address some of these problems. As we will see in the context of our exarnple, they cannot provide a complete solution, but they can provide sorne help. In this chapter, we do not discuss the use of *null* values beyond this one exarnple.

Consider the example Hourly_Elnps relation. Clearly, *null* values cannot help eliminate redundant storage or update anomalies. It appears that they can address insertion and deletion anomalies. For instance, to deal with the insertion anolnaly exarnple, we can insert an elTIplayee tuple with *null* values in the hourly wage field. However, *null* values cannot address all insertion anornalies. For exarnple, we cannot record the hourly wage for a rating unless there is an ernployee with that rating, because we cannot store a null value in the *ssn* field, which is a prirnary key field. Sinlilarly, to deal with the deletion anomaly exarnple, we rnight consider storing a tuple with *null* values in all fields except *rating* and *hourly_wages* if the last tuple with a given *rating* would otherwise be deleted. However, this solution does not work because it requires the *8871,* value to be *null,* and prirnary key fields cannot be *null.* Thus, *null* values do not provide a general solution to the problerns of reclundancy, even though they can help in sorne cases.

19.1.2 Decompositions

Intuitively, redundancy arises when a relational schcrna forces an association between attributes that is not natural. Functional dependencies (and, for that matter, other Ies) can 'be used to identify such situations and suggest re£1nernents to the schema. The essential idea is that rnany problerns arising fr0111 redundancy can be addressed by replacing a relation 'with a collection of 'smaller' relations.

A. decomposition of a relation schema R consists of replacing the relation scherna by t\vo (or 1no1'e) relation schcrnas that each contain a subset of the attributes of R and together include all attributes in R. Intuitively, we want to store the inforrnation in any given instance of R by storing projections of the instance. This section exalnines the use of decornpositions through several exanlples.

We can decornpose lIourly_Ernps into two relations:

Hourly_Emps2(*ssn*, *naTne*, *lot*, *rating*, *hours_worked*)
\Vages(*rating*, *hourly_wages*)

The instances of these relations corresponding to the instance of Hourly_Emps relation in Figure 19.1 is shown in Figure 19.2.

ssn	narne	lot	rating	hours_worked
123-22-3666	Attishoo	48	8	40
231-31-5368	Sluiley	22	8	30
131-24-3650	Smethurst	35	5	30
434-26-3751	Guldu	35	5	32
612-67-4134	Madayan	35	8	40

rating	hourly_wages
8	10
5	7

Figure 19.2 Instances of Hourly_Emps2 and vVages

Note that we can easily record the hourly wage for any rating sirnply by adding a tuple to Wages, even if no ernployee with that rating appears in the current instance of flourly_Ernps. Changing the wage associated \vith a rating involves updating a single Wages tuple. This is rnore efficient than updating several tuples (as in the original design), and it elirninates the potential for inconsistency.

19.1.3 Problems Related to Decomposition

lJnless we are careful, decornposing a relation scherna can create 1n01'e problerns than it solves. Two irnportant questions llHlst be asked repeatedly:

1. 1)0 we need to decornpose a relation?

2. What problerns (if any) does a given deeornposition cause?

To help with the first question, several *normal forms* have been proposed for relations. If a relation scherna is ill one of these nOfrual 1'orms, we know that certain kinds of problerlls cannot arise. Considering the norrnal forrn of a given relation scherna can help us to decide \vhether or not to decornpose it further. If we decide that a relation scherna 111USt be decornposed further, we rnust choose a particular dec()lnposition (I.e., a particular collection of slnaller relations to replace the given relation).

With respect to the second question, two properties of decornpositions are of particular interest. The *lossless-join* property enables us to recover any instance of the decornposed relation froln corresponding instances of the s111aller relations. The *dependency-preservation* property enables us to enforce any constraint on the original relation by sinlply enforcing SaIne contraints on each of the srnaller relations. That is, we need not perform joins of the slllaller relations to check whether a constraint on the original relation is violated.

From a performance standpoint, queries over the original relation may require us to join the decomposed relations. If such queries are common, the perfor-rnance penalty of decomposing the relation may not be acceptable. In this case, we may choose to live with some of the problems of redundancy and not decompose the relation. It is important to be aware of the potential problerns caused by such residual redundancy in the design and to take steps to avoid thern (e.g., by adding SaIne checks to application code). In sonle situations, decomposition could actually *improve* performance. This happens, for exam-ple, if lnost queries and updates exanline only one of the decornposed relations, which is srnaller than the original relation. We do not discuss the irnpact of decompositions on query perforInance in this chapter; this issue is covered in Section 20.8.

()ur goal in this chapter is to explain S0111e powerful concepts and design guide-lines based on the theory of functional dependencies. A good database designer should have a firm grasp of nor1nal fonns and \vhat problerns they (do or do not) alleviate, the technique of decornposition, and potential problerns with decornpositions. For example, a designer often asks questions such as these: Is a relation in a given nonnal forIn? Is a decornposition clependency-preserving? Our objective is to explain when to raise these questions and the significance of the answers.

19.2 FUNCTIONAL DEPENDENCIES

A functional dependency (FD) is a kind of Ie that generalizes the concept of a *key*. Let *R* be a relation scherna and let *X* and Y be nonernpty sets of attributes in *R*. We say that an instance *r* of *R* satisfies the FD$X \rightarrow Y$ [1] if the following holds for every pair of tuples tl and *t2* in *r*:

If $t1.X = t2..X$, then $tl.\}T = t2.Y$.

We use the notation tl.X to refer to the projection of tuple t1 onto the attributes in *X*, in a natural extension of our TRC notation (see Chapter 4) *t.a* for referring to attribute *a* of tuple t. An FD $X \rightarrow$ Yessentially says that if two tuples agree on the values in attributes *X*, they 111Ust also agree on the values in attributes *Y*.

Figure 19.3 illustrates the rneaning of the FD $AB \rightarrow$ C by showing an instance that satisfies this dependency. The first two tuples show that an FD is not the same as a key constraint: Although the FD is not violated, *AB* is clearly not a key for the relation. The third and fourth tuples illustrate that if two tuples differ in either the *A* field or the *B* field, they can differ in the C field without violating the FD. On the other hand, if we add a tuple (aI, bl, *c2, dl)* to the instance shown in this figure, the resulting instance would violate the FD; to see this violation, compare the first tuple in the figure with the new tuple.

A	*B*	*C*	*D*
a1	b1	c1	d1
a1	b1	c1	d2
a1	b2	c2	dl
a2	bl	c3	ell

Figure 19.3 An Instance that Satisfies $AB \rightarrow C$

Recall that a *legal* instance of a relation nUlst satisfy all specified Ies, including all specified FDs. As noted in Section 3.2, Ies rIlust be identified and specified based on the sernantics of the real-world enterprise being n1odeled. By looking at an instance of a relation, we rnight be able to tell that a certain FD does *not* hold. I-Iowever; we can never deduce that an FD *docs* hold by looking at one or 1I10re instances of the relation, beca,use an FD, like other Ies, is a staternent about *all* possible legal instances of the relation.

[1] $X \rightarrow Y$ is read as X *functionally determines Y*, or simply as X *determines Y*.

.A prirnary key constraint is a special case of an FD. The attributes in the key play the role of X, and the set of all attributes in the relation plays the role of Y. Note, however, that the definition of an FD does not require that the set X be 111inilnal; the additionalrninimality condition Illust be Inet for X to be a key. If $X \rightarrow Y$ holds, \vhere Y is the set of all attributes, and there is SCHne (strictly contajned) subset V of X such that $V \rightarrow Y$ holds, then X is a *superkey*.

In the rest of this chapter, ·we see several exarIlples of FDs that are not key constraints.

19.3 REASONING ABOUT FDS

Given a set of FDs over a relation scheula .R, typically several additional FDs hold over R whenever all of the given FDs hold. As an exalnple, consider:

Workers(*ssn, naTne, lot, did, since)*

We know that *ssn* $\rightarrow$ *did* holds, since *ssn* is the key, and FD *did* $\rightarrow$ *lot* is given to hold. Therefore, in any legal instance of Workers, if two tuples have the same *ssn* value, they Blust have the sarne *did* value (frolH the first FD), and because they have the sarrle *did* value, they must also have the saIne *lot* value (1'1'0111 the second FD). Therefore, the FD *ssn* $\rightarrow$ *lot* also holds on Workers.

We say that an FD f is implied by a given set F of FDs if f holds on every relation instance that satisfies all dependencies in F; that is, f holds whenever all FDs in F hold. Note that it is not sufficient for f to hold on SaIne instance that satisfies all dependencies in F; rather, f rnust hold on *every* instance that satisfies all dependencies in F.

19.3.1 Closure of a Set of FDs

The set of all .FDs irnplied by a given set F of FDs is called the closllre of F, denoted as F^+. An irnportant question is how we can infer, or cornpute, the closure of a given set F of FDs. The answer is sirnple and elegant. The following three rules, called Armstrong's Axioms, can be applied repeatedly to infer all FI)s irnplied by a set F of FDs. We use X, Y, and Z to denote *sets* of attributes over a relation scherna R:

- Reflexivity: If $X \supseteq Y$, then $X \rightarrow Y$.
- Augn1.entation: If $X \rightarrow Y$, then $XZ \rightarrow YZ$ for any Z.
- Transitivity: If $X \rightarrow Y$ and $Y \rightarrow Z$, then $X \rightarrow Z$.

Theorem 1 *Armstrong's Axioms are* **sound**, *in that they generate only FDs in* F^+ *when applied to a set* F *of FDs. They are also* **complete**, *in that repeated a]Jplicat'ion afthese rules will generate all FDs in the closure* .Fl+.

The soundness of Arrnstrong's Axiorns is straightfor\vard to prove. Cornpleteness is harder to show; see Exercise 19.17.

It is convenient to use SOlne additional rules while reasoning about $P+:$

- Union: If $X \rightarrow$ Yand $X \rightarrow Z$, then $X \rightarrow YZ$.
- Decomposition: If $X \rightarrow YZ$, then $X \rightarrow y'$ and $X \rightarrow Z$.

These additional rules are not essential; their soundness can be proved using Armstrong's AxiolllS.

To illustrate the use of these inference rules for FDs, consider a relation schelua ABC with FDs $A \rightarrow B$ and $B \rightarrow$ C. In a trivial FD, the right side contains only attributes that also appear on the left side; such dependencies always hold due to reflexivity. Using reflexivity, we can generate all trivial dependencies, which are of the form:

$X \rightarrow$ Y, where $Y \subseteq X$, $X \subseteq ABC$, and $Y \subseteq ABC$.

FrOHl transitivity we get $A \rightarrow$ C. Fronl auglnentation we get the nontrivial dependencies:

$AC \rightarrow BC$, $AB \rightarrow AC$, $AB \rightarrow C13$.

As another exalnple, we use a rnore elaborate version of Contracts:

Contracts(*contractid, supplierid, projectid, deptid, partid, qty, val'ue)*

\Ve denote the schenla for Contracts as $CSJDPQV$. The rneaning of a tuple is that the contract with *contractid* C is an agreement that supplier S (*supplierid*) 'will supply Q iterns of part? *(par-tid)* to project J (*projectid*) associated with departrnent D (*deptid*); the value V of this contract is equal to *value*.

The following res are known to hold:

1. The contract id C is a key: C $\rightarrow$ *CSJDP(JV.*
2. A project purchases a given part using a single contract: *.II)* $\rightarrow C$.

3. A departInent purchases at most one part froul a supplier: 8D → *P*.

Several additional FDs hold in the closure of the set of given FDs:

From *.IP → C, C → CSJDPQ'V,* and transitivity, we infer *.IP → CSJDPQ V.*

FraIn 8D _→ *P* and augnlentation, we infer *SDJ → JP*.

FraIn *8DJ → .IP, JP → CSJDPQV*, and transitivity, we infer *SDJ → CSJD-PQ V.* (Incidentally, while it Illay appear tenlpting to do so, we *cannot* conclude *SD → CSDPQ V,* canceling *.I* on both sides. FD inference is not like aritlunetic Illultiplication!)

We can infer several additionalFDs that are in the closure by using augruen-tation or decomposition. For exarnple, from *C → CSJDPQ V,* using decompo-sition, we can infer:

C → C, C → 5, C → *J*, C → D, and so forth

Finally, we have a number of trivial FDs from the reflexivity rule.

19.3.2 Attribute Closure

If we just want to check whether a given dependency, say, *X → Y,* is in the closure of a set ***F*** of FDs, we can do so efficiently without cornputing Fl+. We first cornpute the attribute closure X+with respect to *F,* \vhich is the set of attributes *A* such that *X → A* can be inferred using the Arrnstrong Axioms. The algorithrn for computing the attribute closure of a set *X* of attributes is shown in Figure 19.4.

```
closure = X;
repeat until there is no change: {
        if there is an FD U → V in F such that U ⊆ closure,
            then set closure = closure ∪ V.
        }
```

Figure 19.4 Computing the Attribute Closure of Attribute Sct *X*

Theorem 2 *The algorithln shown in Figure* 1.9.4 *cornputes the attribute closure X+- of the attribute set X 'IDith respect to the sct of FDs Fl.*

The proof of this theorern is considered in Exercise 19.15. This algoriUuIl can be rl10dified to find keys by starting with set X containing a, single attribute and stopping as soon as *closure* contains all attributes in the relation scherna. By varying the starting attribute and the order in which the algorithrIl considers FDs, we can obtain all candidate keys.

19.4 NORMAL **FORMS**

Given a relation sche111a, we need to decide whether it is a good design or we need to decornpose it into srnaller relations. Such a decision llUlst be guided by an understanding of what problenls, if any, arise froln the current schelna. To provide such guidance, several normal forms have been proposed. If a relation schelna is in one of these norrnal forIns, we know that certain kinds of problerlls cannot arise.

The nonnal forrns based on FDs are *first nor-rnal forrn* (1*NF)*, *second normal forrn (2NF)*, *third normal form (3NF)*, and *Boyce-Codd normal form (BCNF).* These fonns have increasingly restrictive requirernents: Every relation in BCNF is also in 3NF, every relation in 3NF is also in 2NF, and every relation in 2NF is in INF. A relation is in first normal fortH if every field contains only atornic values, that is, no lists or sets. This requirerllent is iInplicit in our definition of the relational rnode!. Although SOHle of the newer database systerlls are relaxing this requirernent, in this chapter we aSSUlne that it always holds. 2NF is Inainly of historical interest. 3NF and BCNF are irnportant frolH a database design standpoint.

While studying norrnal fonns, it is irnportant to appreciate the role played by FDs. Consider a relation scherna R with attributes ABC. In the absence of any ICs, any set of ternary tuples is a legal instance and there is no potential for redundancy. ()n the other hand, suppose that we have the FI) $A \rightarrow 13$. Now if several tuples have the sarne A value, they rnust also have tllC sarneB value. This potential redundancy can be predicted using the FD il1fonnation. If 11101'8 detailed 1Cs are specified, we rnay be able to detect rnore subtle redundancies as well.

We primarily discuss redundancy revealed l)y PI) inforrnation. In Section 19.8, we discuss 11lore sophisticated 1Cs called *multivalued dependencies* and *join dependencies* and norrnal forrns based on theIn.

19.4.1 Boyce_Codd Normal Form

Let R be a relation scherna, F be the set ofF'I)s given to hold over R, X be a subset of the attributes ofR, and A be an attribute of R. R is in Boyce-Codd

normal form if, for everyFl)$X \rightarrow A$ in F, one of the follo\ving statements is true:

- $A \in X$; that is, it is a trivial FD, or
- X is a superkey.

Intuitively, in a BCNF relation, the only nontrivial dependencies are those in 'which a key detennines SaIne attribute(s). Therefore, each tuple can be thought of as an entity or relationship, identified by a key and described by the reluaining attributes. !(ent (in [425]) puts this colorfully, if a little loosely: "Each attribute nlust describe [an entity or relationship identified by] the key, the \vhole key, and nothing but the key." If we use ovals to denote attributes or sets of attributes and draw arcs to indicate FDs, a relation in BCNF has the structure illustrated in Figure 19.5, considering just one key for simplicity. (If there are several candidate keys, each candidate key can play the role of KEY in the figure, with the other attributes being the ones not in the chosen candidate key.)

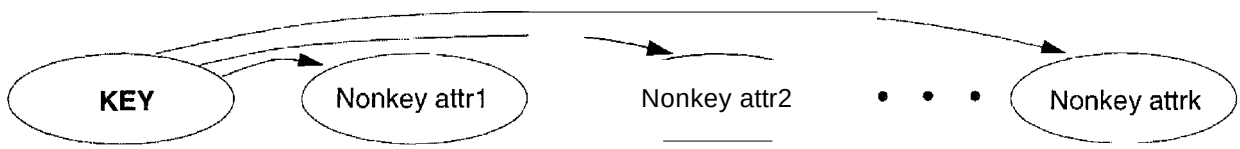


Figure 19.5 FDs in a BCNF Relation

BCNF ensures that no redundancy can be detected using FD infonnation alone. It is thus the Inost desirable norrnal form (fronl the point of view of redundancy) if we take into account only FD information. This point is illustrated in Figure 19.6.

X	Y	A
x	y_1	a
x	y_2	?

Figure 19.6 Instance Illustrating BCNF

This figure shows (t\VO tuples in) an instance of a relation with three attributes X, Y, and A. There a.re two tuples with the saIne value in the X colurnn. Now suppose that we kno\v that this instance satisfies an FD $X \rightarrow A$. We can see that one of the tuples has the value a in the A colurnn. What can we infer al)out the value in the A colllrnn in the second tuple? 'Using the FI), \ve can conclude that the second tuple also has the value a in this colurnn. (Note that this is really the only kind of inference we can make about values in the fields of tuples by usingFDs.)

But is this situation not an exaInple of redundancy? \Ve appear to have stored the value a twice. Can such a situation arise in a BCNF relation? The ans\ver is No! If this relation is in BCNF, because A is distinct fronl X, it follows that X IllU8t be a key. (Otherwise, the FD $X \rightarrow A$ \vould violate BCNF.) If X is a key, then Yl = $Y2$, which Ineans that the two tuples are identical Since a relation is defined to be a *set* of tuples, we cannot have two copies of the saIne tuple and the situation shc)\vn in Figure 19.6 cannot arise.

rrherefore, if a relation is in BCNF, every field of every tuple records a piece of inforlnation that cannot be inferred (using only FDs) frorn the values in all other fields in (all tuples of) the relation instance.

19.4.2 Third Normal Form

Let R be a relation scherna, F be the set of FDs given to hold over R, X be a subset of the attributes of R, and A be an attribute of R. R is in third normal forIn if, for every FD $X \rightarrow A$ in F, one of the following statenlents is true:

- $A \in X$; that is, it is a trivial FD, or
- X is a superkey, or
- A is part of sorne key for R.

rrhe definition of 3NF is sinlilar to that of BCNF, with the only difference being the third condition. Every BCNF relation is also in 3NF. To understand the third condition, recall that a key for a rela,tion is a *minimal* set of attributes that uniquely deterrnines all other attributes. A rrlllst be part of a key (any key, if there are several). It is not enough for A to be part of a superkey, because the latter condition is satisfied by every attribute! Finding all keys of a relation scherna is known to be an NP-cornplete problern, and so is the prob1ern of detennining whether a relation seherna is in 3NF.

Suppose that a dependency $X \rightarrow A$ causes a violation of 3NF. There are two cases:

- X is *a proper subset of some key* K. Such a dependency is 801netirnes called a partial dependency. In this case, we store *(X, ./1)* pairs redundantl:y. As an example, consider the Reserves relation with attributes $SBDC$ frorn Section 19.7.4. The only key is *8El)*, and we have the FD $S \rightarrow C$. We store the credit ca,rd nurnber for a sailor as lnany tirnes as there are reservations for that sailor.
- X is *not a proper subset of any key.* Such a dependerlcy is sornetirnes called a transitive **dependency**, because it rneans we have a chain of

dependencies $!(\rightarrow X \rightarrow A$. The problem is that we cannot associate an X value with a K value unless we also associate an A value with an X value. As an exanlple, consider the Hourly_Emps relation with attributes *SNLRWH* froIn Section 19.7.1. The only key is *S*, but there is an FD $R \rightarrow W$, which gives rise to the chain $S \rightarrow R \rightarrow W$. The consequence is that we cannot record the fact that elnployee *S* has rating *R* without knowing the hourly \vage for that rating. This condition leads to insertion, deletion, and update anoIllalies.

Partial dependencies are illustrated in Figure 19.7, and transitive dependencies are illustrated in Figure 19.8. Note that in Figure 19.8, the set X of attributes 11lay or Illay not have some attributes in conunon with KE-Y; the diagranl should be interpreted as indicating only that X is not a subset of KEY.

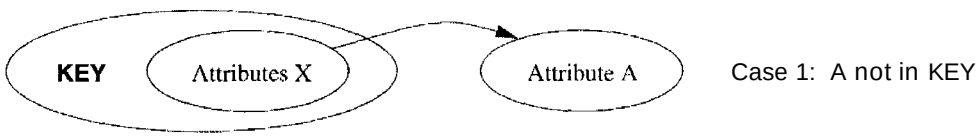


Figure 19.7 Partial Dependencies

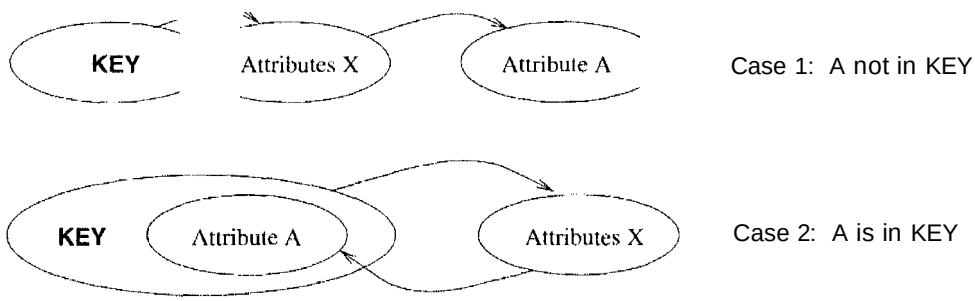


Figure 19.8 Transitive Dependencies

The Inotivation for 3NF is rather technical. By Inaking an exception for certain dependencies involving key attributes, we can ensure that every relation schclna can be decornposed into a collection of 3NF relations using only dec(nnpositions that have certain desirable properties (Section 19.5). Such a guarantee does not exist for BCNF relations; the 3NF definition weakens the BCNF requirernents just enough to Inake this guarantee possible. We Inay therefore cOlnprornise by settling for a 3NF design. As we see in Chapter 20, we 11lay sometimes accept this compromise (or even settle for a non-3NF scheIna) for other reasons as well.

lJnlike BCNF, however, BOlne redundancy is possible with 3NF. The problerns associated \vith partial and transitive dependencies persist if there is a nontrivial dependency $X \rightarrow A$ and X is not a superkey, even if the relation is in 3NF l)ccause *A* is pa,rt of a key. To understand this point, let us revisit the R,eserves

relation with attributes *SEDe* and the FD $S \rightarrow C$, which states that a sailor uses a unique credit card to pay for reservations. *S* is not a key, and *C* is not part of a key. (In fact, the only key is *SED.)* Hence, this relation is not in 3NF; *(S, CJ* pairs are stored redundantly. IIowever, if we also know that credit cards uniquely identify the owner, we have the FD C → 5, which rneans that *CBD* is also a key for Reserves. Therefore, the dependency $S \rightarrow C$ does not violate 3NF, and R,eserves is in 3NF. Nonetheless, in all tuples containing the saIne 5 value, the saIne (8, *CJ* pair is redundantly recorded.

For cOlllpleteness, we reluark that the definition of second norrnal form is essentially that partial dependencies are not allowed. Thus, if a relation is in 3NF (which precludes both partial and transitive dependencies), it is also in 2NF.

19.5 PROPERTIES OF DECOMPOSITIONS

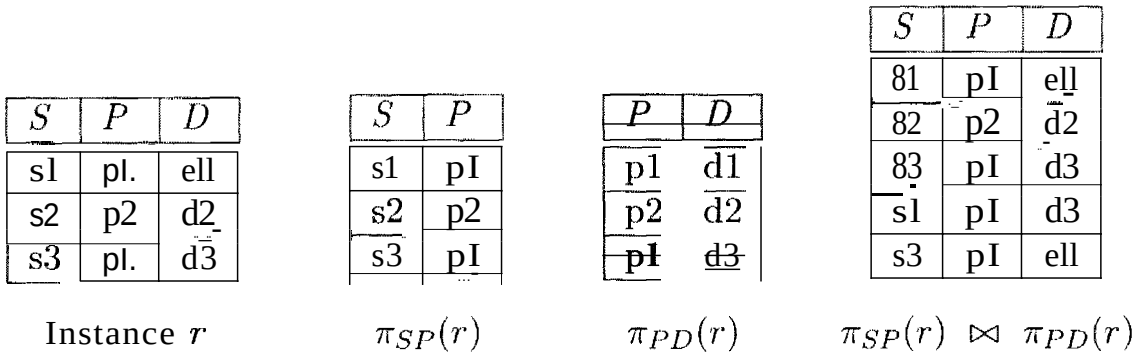
DecoIllposition is a tool that allows us to eliminate redundancy. As noted in Section 19.1.3, however, it is important to check that a decoInposition does not introduce new problellls. In particular, we should check whether a decomposition allows us to recover the original relation, and whether it allows us to check integrity constraints efficiently. We discuss these properties next.

19.5.1 Lossless-Join Decomposition

Let *R* be a relation schelna and let *F* be H, set of FDs over *R.* A decolnposition of *R* into two schernas with attribute sets *X* and Y is said to be a lossless-join decomposition with respect to *F* if, for every instance *r* of *R* that satisfies the dependencies in *F*, $\pi_X(r) \bowtie \pi_Y(r) = r$. In other words, we can recover the original relation 1'rom the deconlposed relations.

This definition can easily be extended to cover a decornposition of *R* into more than two relations. It is easy to see that $r \subseteq \pi_X(r) \bowtie \pi_Y(r)$ ahvays holds. III general, though, the other direction does not hold. If we take projections of a relation and recornbine theln using natural join, we typically obta.in SOIne tuples that 'were not in the original relation. This situation is illustrated in Figure 19.9.

By replacing the instance *r* shown in Figure 19.9 with the instances $\pi_{SP}(r)$ and $\pi_{PI)}(r)$, we lose sorne information. In particular, suppose that the tuples in *r* denote relationships. We can no longer tell that the relationships $(81, p_1, d_3)$ and (s_3, p_1, d_1) do not hold. rrhe decoluposition of schema *SPD* into *S.P* and *PI)* is therefore loss,Y if the instance *r* shown in the figure is legal, that is, if this



Instance r

S	P	D
s1	pl.	ell
s2	p2	d2
s3	pl.	d3

$\pi_{SP}(r)$

S	P
s1	pI
s2	p2
s3	pI

$\pi_{PD}(r)$

P	D
p1	d1
p2	d2
pI	d3

$\pi_{SP}(r) \bowtie \pi_{PD}(r)$

S	P	D
81	pI	ell
82	p2	d2
83	pI	d3
s1	pI	d3
s3	pI	ell

Figure 19.9 Instances Illustrating Lossy Decompositions

instance could arise in the enterprise being rIlodeled. (Observe the siInilarities between this example and the Contracts relationship set in Section 2.5.3.)

All decompositions used to eli'minate redundancy **must** *be lossless.* The following sirnple test is very useful:

Theorern 3 *Let* R *be a relation and* F *be a set of FDs that hold over* R. *The decomposition of* R *into relations with attribute sets* R_1 *and* R_2 *is l08sless if and only if* F^+ *contains either the FD* $R_1 \cap R2 \rightarrow R_1$ *or the FD* $R_1 \cap R_2 \rightarrow R2$.

In other words, the attributes cornrIlon to *Rl* and *R2* HUlst contain a key for either RIOI' R_2.[2] If a relation is decornposed into 1110re than two relations, an efficient (time polynomial in the size of the dependency set) algoritllln is available to test whether or not the dec(nnposition is lossless, but we will not discuss it.

Consider the lIourly_Ernps relation again. It has attributes *SNLRWII,* and the FI) $R \rightarrow W$ causes a violation of 3NF. We dealt with this violation by decorIlposing the relation into *SNLRII* and RW. Since R is cornrnon to both decornposed relations and $R \rightarrow W$ holds, this decornposition is lossless-join.

This exarnple illustrates a general observation that follows froIH Theorerll 3:

> If an Ff) $X \rightarrow Y$ holds over a relation R and $X \cap Y$ is ernpty, the decornposition ofR into $R - Y$ and XY is lossless.

X appears in both $R - Y$ (since $X \cap Y$ is ernpty) and XY, and it is a key for XY.

[2]See Exercise 19.19 for a proof of Theorern 3. Exercise 19.11 illustrates that the 'only if' claim depends on the assumption that only functional dependencies can be specified as integrity constraints.

Another hnportant observation, which we state without proof, has to do with repeated decolnpositiollS. Suppose that a relation R is decornposed into Rl and $R2$ through a IOBsless-join decolupositiol1, and that $R1$ is decolnposed *intoRl.1* and $R12$ through another lossless-join decolnposition. Then, the decolnposition of R into $R11$, $R.12$, and $R2$ is lossless-join; by joining $R11$ and $R12$, \ve can recover $R.1$, and by then joining RI and $R2$, we can recover R.

19.5.2 Dependency-Preserving Decomposition

Consider the Contracts relation with attributes *C8JDPCJVfronl* Section 19.3.l. The given FDs are C $\rightarrow$ *C8JDPQV*, *JP* $\rightarrow$ *C*, and *SD* $\rightarrow$ *P*. Because *SD* is not a key the dependency *SD* $\rightarrow$ *P* causes a violation of BCNF.

We can decolnpose Contracts into two relations with schelnas *CSJDQV* and *SDP* to address this violation; the decolnposition is lossless-join. There is one subtle problelll, however. We can enforce the integrity constraint *JP* $\rightarrow$ C easily when a tuple is inserted into Contracts by ensuring that no existing tuple has the same *JP* values (as the inserted tuple) but different *C* values. Once we decompose Contracts into *CSJDQV* and *SDP*, enforcing this constraint requires an expensive join of the two relations whenever a tuple is inserted into *CSJDQV*. We say that this decornposition is not dependency-preserving.

Intuitively, a *dependency-preserving decornposition* allows us to enforce all FDs by exarnining a single relation instance on each insertion or rnodification of a tuple. (Note that deletions cannot cause violation of FDs.) To define dependency-preserving decornpositions precisely, we have to introduce the concept of a projection of FDs.

Let R be a relation schenla that is decolnposed into two schemas with attribute sets X' and Y, and let F be a set of FDs over R. The **projection of F on X** is the set of FDs in the closure $I^{i'}+$ (not just $.F$!) that involve only attributes in X. We denote the projection of F on attributes X as F_X. Note that a dependency $U \rightarrow V$ in $F+$ is in F_X only if *all* the attributes in [/ and V are in X.

The decornposition of relation scherna R with FI)s F into schcrnas with attribute sets X and Y is dependency-preserving if $(F_X \cup F_Y)^+ = F^+$. That is, if we take the dependencies in F_X and F_Y and cornpute the closure of their un.ion, we get back all dependencies in the closure of F. rrherefore, we need to enforce onl,y the dependencies in F_X and F_Y; allFDs in F^+ are then sure to be satisfied. To enforce F_X,\vs need to examine only relation)((on in.serts to that relation). To enforce F_Y, we need to exarnine only relation Y.

To appreciate the need to consider the closure F^+ while COIUpllting the projection of F, suppose that a relation R with attributes ABC is decomposed into relations\vith attributes AB and $Be:$. The set F of FDs overR includes $A \rightarrow B$, $B \rightarrow C$, and $C \rightarrow A$. Of these, $A \rightarrow B$ is in F_{AB} and $B \rightarrow C$ is in F_{BC}. But is this decoIIlposition dependency-preserving? What about $C \rightarrow A$? This dependency is not irnplied by the dependencies listed (thus far) for F_{AB} and F_{BC}.

The closure of F contains all dependencies in F plus $A \rightarrow C$, $B \rightarrow A$, and $C \rightarrow B$. Consequently, F_{AB} also contains $B \rightarrow A$, and F_{BC} contains $C \rightarrow B$. Therefore, $F_{AB} \cup F_{BC}$ contains $A \rightarrow B$, $B \rightarrow C$, $B \rightarrow A$, and $C \rightarrow B$. The closure of the dependencies in F_{AB} and F_{BC} now includes $C \rightarrow A$ (which follows frorn $C \rightarrow B$, $B \rightarrow A$, and transitivity). l'hus, the deccHnposition preserves the dependency $C \rightarrow A$.

A direct application of the definition gives us a straightforward algoritlun for testing whether a deconlposition is dependency-preserving. (This algorithrn is exponential in the size of the dependency set. A polynomial algorithnl is available; see Exercise 19.9.)

We began this sectiol1with an exanlple of a lossless-join deC0111position that was not dependency-preserving. Other decorupositions are dependency-preserving, but not lossless. A silnple example consists of a relation ABC' with FD $A \rightarrow B$ that is decornposed into AB and BG.

19.6 NORMALIZATION

Having covered the concepts needed to understand the role of HortHa} forms and decolnpositions in database design, we now consider algoritlnIls for converting relations to BCNF or 3NF. If a relation schema is not in BCNF, it is possible to obtain a lossless-join deccunpositioll into a collection of BCNF relation schemas. Unfortunately, there may be no dependenc,y-preserving decOlnposition into a collection of BCN.F relation schernas. However, there is always a dependency-preserving, lossless-join decoruposition into a collection of 3NF relation schernas.

19.6.1 Decomposition into BCNF

We now present an algorithm for decornposing a relation scherna R with a set of FI)sF into a collection of BCNF relation schernas:

1. Suppose that R is not in BCNF. Let $X \subset R$, A be a single attribute in R, and $X \rightarrow A$ be an FD that causes a violation of BCNF. DecornposeR into $R - A$ and XA.
2. If either $R - A$ or XA is not in BCN.F, decornpose thern further by a recursive application of this algorithrn.

$R - A$ denotes the set of attributes other than A in R, and XA denotes the union of attributes in X and A. Since $X \rightarrow A$ violates BCNF, it is not a trivial dependency; further, A is a single attribute. Therefore, A is not in X; that is, $X \cap A$ is ernpty. Therefore, each dec()lnposition carried out in Step 1 is lossless-join.

The set of dependencies associated with $R - A$ and XA is the projection of F onto their attributes. If one of the new relations is not in BCNF, we decornpose it further in Step 2. Since a decornposition results in relations with strictly fewer attributes, this process terrninates, leaving us with a collection of relation schernas that are all in BCNF. Further, joining instances of the (two or 1nore) relations obtained through this algorithrn yields precisely the corresponding instance of the original relation (i.e., the decorllposition into a collection of relations each of which in BCNF is a lossless-join dec()lnposition).

Consider the Contracts relation with attributes $C3JDPQV$ and key C. We are given FDs $JP \rightarrow C$ and $3D \rightarrow P$. By using the dependency $3D \rightarrow P$ to guide the decornposition, we get the two schernas $3DP$ and $C5JDQV$. $51)P$ is in BCNF. Suppose that we also have the constraint that each project deals with a single supplier: $.I \rightarrow 5$. This rneans that the sche1na $CSJDQV$ is not in BCNF So we deccnnpose it further into $J3$ and $C.IDC2V$. $C \rightarrow JDQV$ holds over $CJDQV$; the only other FI)s that hold are those obtained frorll this PI) by augrnentation, and therefore all FDs contain a key in the left side. Thus, each of the schernas $ST)P$, $.IS$, and $CJDQV$ is in BCNF, and this collection of schcrnas also represents a lossless-join decornposition of $CSJDQV$.

The steps in this deC(nllposition process can be visualized as a tree, as shown in Figure 19.10. The root is the original relation $CSJIJPQV$, and the leaves are the BCNF relations that result frorn the deccHnposition aJgorithrn: $3D?$, $.IS$, and $CSDQV$. Intuitively, ea.ch internal node is replaced by its children through a single decomposition step guided by the FD shown just below the node.

Redundancy in BCNF Revisited

The decolnposition of $CSJDQV$ iuto SDP, JS, and $CJDQV$ is not dependency-preserving. Intuitively, dependency $Jp \rightarrow C$ carlllot be enforced without a, join. ()ne way to deal \vith this situation is to add a relation \vith attributes $GJ\}$. In

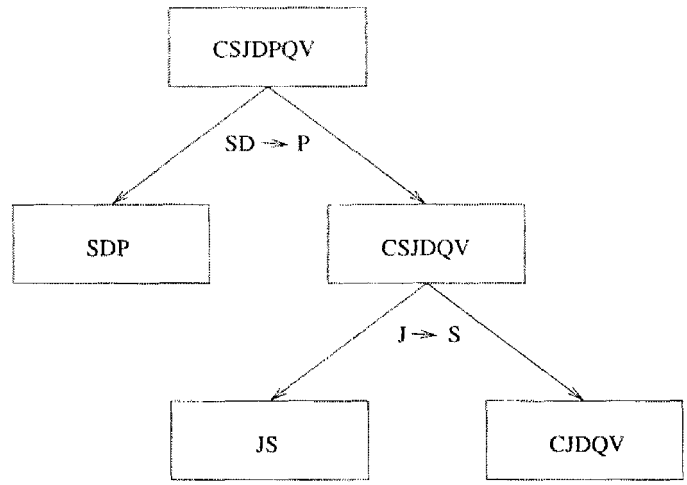


Figure 19.10 Decomposition of *CSJDQV* into *SDP*, *JS*, and *CJDQV*

effect, this solution arnounts to storing SOITle information redundantly to rnake the dependency enforcement cheaper.

This is a subtle point: Each of the schemas *CJP*, *SDP*, *JS*, and *CJDQV* is in BCNF, yet some redundancy can be predicted by FD infonnation. In particular, if we join the relation instances for *SDP* and *CJDQVand* project the result onto the attributes *CJP*, we rnust get exactly the instance stored in the relation with scherna *CJP*. We saw in Section 19.4.1 that there is no such redundancy within a single BCNF relation. This exarnple shows that redundancy can still occur across relations, even though there is no redundancy within a relation.

Alternatives in Decomposing to BCNF

Suppose several dependencies violate BCNF. Depending on ·which of these dependencies we choose to guide the next decornposition step, we rnay arrive at quite different collections of BeNF relations. Consider Contracts. We just decornposed it into *SDP*, *is*, and *CJDQV*. Suppose we choose to decornpose the original relation *CSJDPQV* into *JS* and *CJDPQV*, based on the FD *J* → *S*. The only dependencies that hold over *CJDPQV* are *JP* → C and the key dependency C → *CJDPQV*. Since *JP* is a key, *CJDPQV* is in BeNF. Thus, the schernas *JS* and *CJDPQV* represent a lossless-join decornposition of Contracts into BCNF relations.

The lesson to be learned here is that the theor,Y of dependencies can tell us ·when there is redundancy and give us clues about possible clecornpositions to address the problern, but it cannot discrirninate arnong decornposition alternatives. A designer has to consider the alternatives and choose one based on the scrnantics of the application.

BCNF and Dependency-Preservation

Sometimes, there siIuply is no decomposition into BCNF that is dependency-preserving. As an exaruple, consider the relation schelna *SBD,* in which a tuple denotes that sailor *S* has reserved boat ,8 Oll date *[J.* If we have the FDs *8B* $\rightarrow$ *D* (a sailor can reserve a given boat for at nlost one day) and D $\rightarrow$ *B* (on any given day at rllost one boat can be reserved), *SBn* is not in BCNF because *D* is not a key. If we try to dec(nnpose it, however, we cannot preserve the dependency *BB* $\rightarrow$ *D.*

19.6.2 Decomposition into 3NF

Clearly, the approach we outlined for 10ssless-joiIl decornpositioIl into BCNF also gives us a lossless-join decomposition into 3NF. (Typically, we can stop a little earlier if we are satisfied with a collection of 3NF relations.) But this approach does not ensure dependency-preservation.

A siInple rllodification, however, yields a deco111position into 3NF relations that is lossless-join and dependency-preserving. Before we describe this modification, we need to introduce the concept of a lninirnal cover for a set of FDs.

Minimal Cover for a Set of FDs

A minimal cover for a set *F* of FDs is a set G of FDs such that:

1. Every dependency in *G* is of the forIn $X \rightarrow A$, where *A* is a single attribute.
2. The closure *F+* is equal to the closure (;+.
3. If we obtain a set *II* of dependencies frorn G by deleting one or 1110re dependencies or by deleting attributes frorn a dependency in G, then $p+' \neq II+$.

Intuitively, a rninirnal cover for a set *F* of FDs is an equivalent set of dependencies that is *minimal* in two respects: (1) Every dependency is as slIlall as possible; tha,t; is, each attribute on the left side is necessary and the right side is a single attribute. (2) Every dependency in it is required for the closure to be equal to F^+.

As an exarnplc, let *F* be the set of dependencies:

$$\text{it} \rightarrow B,\ ABCD \rightarrow E,\ EF \rightarrow G,\ F \rightarrow I,\ \text{and}\ CDF \rightarrow EG.$$

First, let us rewrite it$CDF \rightarrow BG$ so that every right side is a single attribute:

$ACDF \rightarrow E$ and $ACDF \rightarrow G$,

Next consider $ACDF \rightarrow G$, This dependency is irnplied by the following FDs:

$A \rightarrow B$, $ABCD \rightarrow E$, and $EF \rightarrow G$,

Therefore, we can delete it, Sirnilarly, we can delete $ACDF \rightarrow E$. Next consider $ABCD \rightarrow E$, Since $A \rightarrow B$ holds, we can replace it with $ACD \rightarrow E$, (At this point, the reader should verify that each rernaining FD is rninilnal and required,) Thus, a minimal cover for F is the set:

$A \rightarrow B$, $ACD \rightarrow E$, $EF \rightarrow G$, and $EF \rightarrow H$,

The preceding exarnple illustrates a general algorithrn for obtaining a rninimal cover of a set F of FDs:

1. Put the FDs in a Standard Form: Obtain a collection G of equivalent FDs with a single attribute on the right side (using the decornposition axiolIl),
2. Minimize the Left Side of Each FD: For each FD in G, check each attribute in the left side to see if it can be deleted while preserving equivalence to F^+,
3. Delete Redundant FDs: Check each reluaining FD in G to see if it can be deleted while preserving equivalence to F^+,

Note that the order in which we consider FDs while applying these steps could produce different rninilnal covers; there could be several rninirnal covers for a given set of FDs,

IV101'8 irnportant, it is necessary to rniniInize the left sides of F'Ds *before* checking for redundant FI)s, If these two steps are reversed, the final set of FI)s could still contain senne redundant FDs (i,e., not be a rninirnal cover), as the following exarnple illustrates, LetF be the set of dependencies, each of which is already in the standard fornl:

$ABCD \rightarrow E$, $E \rightarrow D$, $A \rightarrow B$, and $AC \rightarrow D$,

Observe that none of these FDs is redundant; if we checked for redundantFDs first, we would get the same set of FI)s F. The left side of *il13CIJ* E can be replaced by AC while preserving equivalence to F^+, and we \vould stop here if \ve checked for reclunda.ntF'Ds in F before rnillilnizing the left sides. However, the set of FDs we have is not a Inininlal cover:

$AC \rightarrow E, E \rightarrow D, A \rightarrow B$, and $AC \rightarrow D$.

From transitivity, the first two FDs irnply the last FD, which can therefore be deleted while preserving equivalence to F^+. The irnportant point to note is that $AC \rightarrow D$ becc)lnes redundant only after we replace $ABeD \rightarrow E$ with $AC \rightarrow E$. If we Ininirnize left sides of FDs first and then cheek for redundantFDs, we are left with the first three FDs in the preeeding list,whieh is indeed a Ininirna1 cover for F.

Dependency-Preserving Decomposition into 3NF

Returning to the problenl of obtaining a lossless-join, dependency-preserving decornposition into 3NF relations, let R be a relation with a set [/' of FDs that is a minirnal cover, and let $R_1, R_2, \ldots, R_n$ be a lossless-join decolnposition of R. For $1 \leq i \leq n$, suppose that each R_i is in 3NF and let F_i denote the projection of F onto the attributes of R_i. Do the following:

- Identify the set N of dependencies in F that is not preserved, that is, not included in the closure of the union of *Fis.*
- :F'or each FD $X \rightarrow A$ in N, create a relation schelna XA and add it to the decomposition of R.

Obviously, every dependency in F is preserved if we replace R by the R_is plus the schernas of the forn1 XA added in this step. The *Ris* are given to be in 3NF. We can show that each of the schemas XA is in 3NF as follows: Since $X \rightarrow A$ is in the lninirnal cover F, $Y \rightarrow A$ does not hold for any Y that is a strict subset of X. Therefore, X is a key for XA. :F\llrther, if any other dependencies hold over XA, the right side can involve only attributes in X' because A is a single attribute (because $X \rightarrow A$ is an FD in a rninhnal cover). Since X is a key for ...:*YA*, none of these additional dependencies causes a violation of 3NF (although they rnight cause a violation of BCNF).

As an optilYlization, if the set N contains several FI)swith the saIne left side, say, $X \rightarrow A_1, X \rightarrow A_2, \ldots, X \rightarrow A_n$, we can replace thern \vith a single equivalent FD $X \rightarrow A_1 \ldots A_n$. Therefore, we produce one relation scherna $XA_1 \ldots A_n$, instead of several schernas $XA_1, \ldots, XAn$, \vhich is generally preferable.

Consider the Contracts relation with attrilnltes *CSJDPQV* and FI)s $JP \rightarrow C$, $SD \rightarrow P$. and $J \rightarrow S$. If we decolnpose *CSJDPQV* into *SDIJ* and *CSJDQV*, then *8DP* is in BCNF, but *CSJDQV* is not even in 3NF. So \ve dec.olupose it further into *JS* and *CJDQV*. rrhe relation schemas *SDP*, *.I8*, and *CJDQV* are in 3NF (in fact, in BCNF), and the decoInposition is lossless-join. However,

the dependency $JP \rightarrow C$ is not preserved. This problerIl can be addressed by adding a relation schema CJP to the decornposition.

3NF Synthesis

We assurned that the design process starts with an ER diagraII1, and that our use of FDs is primarily to guide decisions about decolnposition. The algorithill for obtaining a lossless-join, dependency-preserving decornpositiol1 was presented in the previous section fro111 this perspective--------a lossless-join decoruposition into 3NF is straightforward, and the algorithrn addresses dependency-preservation by adding extra relation schcrnas.

An alternative approach, called synthesis, is to take all the attributes over the original relation R and a rnininlal cover F for the FDs that hold over it and add a relation scherna XA to the decomposition of R for each FD $X \rightarrow A$ in F.

The resulting collection of relation schernas is in 3NF and preserves all FDs. If it is not a lossless-join decomposition of R, we can Dlake it so by adding a relation schenla that contains just those attributes that appear in sorne key. This algorithrn gives us a lossless-join, dependency-preserving decornposition into 3NF and has polynornial corIlplexity-----polynornial algorithms are available for coruputing rninirnal covers, and a key can be found in polync)Inial tirHe (even though finding all keys is known to be NP-cornplete). The existence of a polynornial algorithnl for obtaining a lossless-join, dependency-preserving decornposition into 3NF is surprising when we consider that testing whether a given schema is in 3NF is NP-cornplete.

As an exarnple, consider a relation ABC with FI)s $F = \{A \rightarrow B,\ C \rightarrow B\}$. The first step yields the relation scheluas AB and BG. This is not a lossless-join deC0l11position of $AilC$; $AB\ nBC$ is B, and neither $B \rightarrow A$ nor $B \rightarrow C$ is in F^+. If we add a schema AC, we have the lossless-join property as well. Although the collectic)ll of relations AB,BC, and AC is a depenclency-preserving, lossless-join decornposition of ABC, we obtained it through a process of *synthesis*, rather tllan through a process of repeated decornposition. We note that the decoIIIposition produced by the synthesis approa,ch heavily dependends on the rninirnal cover used.

As another example of the synthesis approach, consider the Contracts relation with attributes $CSJDPQV$ and the followving FI)s:

C $CSJDPQV$, $JP \rightarrow C$, $JD \rightarrow P$, and $J \rightarrow S$.

This set of FI)s is not a rninirnal cover, and so we IIIUSt find OI18. We first replace $G \rightarrow CSJDPQV$ with tllcF'I)s:

$C \rightarrow S$, $C \rightarrow J$, $C \rightarrow IJ$, $C \twoheadrightarrow P$, $C \rightarrow Q$, and $C \rightarrow V$.

The FD $C \rightarrow P$ is irnplied by $C \rightarrow S$, $C \twoheadrightarrow D$, and $SD \rightarrow P$; so we can delete it. The FD $C \rightarrow S$ is irnplied by $C \rightarrow J$ and $J \twoheadrightarrow S$; so we can delete it. This leaves us with a rninirnal cover:

$C \rightarrow J$, $C \rightarrow D$, $C \rightarrow Q$, $C \rightarrow V$, $JP \rightarrow C$, $JD \rightarrow P$, and $J \twoheadrightarrow S$.

lJsing the algorithrll for ensuring dependency-preservation, we obtain the relational scherna *CJ, CD, CQ, CV, GIJP, SDP,* and *JB.* We can irnprove this schenla by cornbining relations for which C is the key into *CDJPQV.* In addition, we have *SDP* and *JS* in our decorllposition. Since one of these relations (*CDJPQV*) is a superkey, we are done.

Conlparing this decomposition with that obtained earlier in this section, we find they are quite close, with the only difference being that one of them has *CDJPQV* instead of *CJP* and *CJDQV.* In general, however, there could be significant differences.

19.7 SCHEMA REFINEMENT IN DATABASE DESIGN

We have seen how normalization can eliminate redundancy and discussed several approaches to nonnalizing a relation. We now consider how these ideas are applied in practice.

Database designers typically use a conceptual design rnethodology, such as ER design, to arrive at an initial database design. Given this, the approach of repeated decorllpositions to rectify instances of redundancy is likely to be the rnost natural use of PI)s and nonnalization techniques.

In this section, we Inotivate the need for a schcrna refinernent step follovving ER design. It is natural to ask whether we even need to decornpose relations produced by translating an ER diagranl. Should a good ER design not lead to a collection of relations free of redundancy prob.lerns? Unfortunately, ER design is a c()!nplex, subjective process, and certain constraints are not expressible in tenns of ER diagraJns. The exaruples in this section are intendecl to illustrate why decornposition of relations produced through ER design rnight be necessary.

19.7.1 Constraints on an Entity Set

Consider the Hourly_Emps relation again. rrhe constraint that attribute *ssn* is a key can be expressed as an FI):

$$\{ssn\} \rightarrow \{ssn,\ name,\ lot,\ rating,\ hourly\_wages,\ hours\_worked\}$$

:For brevity, we \vrite this FD as $S \rightarrow SNLRWH$, using a single letter to denote each attribute and ornitting the set braces, but the reader should rernernber that both sides of an FD contain sets of attributes. In addition, the constraint that the *hourly_wages* attribute is deterruined by the *rating* attribute is an FD: $R \rightarrow W$.

As we saw in Section 19.1.1, this FI) led to redundant storage of rating wage associations. *It cannot be expressed in terms of the ER model. Only FDs that determine all attributes of a relation (i.e., key constraints) can be expressed in the ER rnodel.* rrherefore, we could not detect it when we considered Hourly_EIIlPS as an entity set during ER IIlodeling.

We could argue that the problenl with the original design was an artifact of a poor ER design, which could have been avoided by introducing an entity set called Wage_Table (with attributes *rating* and *hourly_wages*) and a relationship set lIas_Wages associating IIourly_.ErIlps and Wage_Table. The point, however, is that we could easily arrive at the original design given the subjective nature of ER rnodeling. Having forInal techniques to identify the problenl with this design and guide us to a, better design is very useful. The value of such techniques cannot be underestirnated when designing large schernas·····-schcrnas with rnore than a hundred tables are not unCOIlllIHon.

19.7.2 Constraints on a Relationship Set

The previous exarnple illustrated how FDs can help to refine the subjective decisions Blade during ER. design, but one could argue that the best possible ER, eliagrarn \vould have led to the same final set of relations. ()ur next exarnple shows how Ff) inforrnation ca.ll lead to a. set of relations unlikely to be arrived at solely through ER design.

We revisit an example frOTH Chapter 2. Suppose that we have entity sets Parts, Suppliers, and I)epartrnents, as \vell as a relationship set Contracts that involves all of theIn. We refer to the scherna for Contra(:ts as $CQPSD$. A contra,ct with contract id C specifies that a supplier S will supply sorne quantity Q of a part P to a departrnent J). (We have adderl the contract iel field C' tC) tIle versiorl of the Contracts relation discussed in Chapter 2.)

We Blight have a policy that a departrnent purchases at Inost one part frorl1 any given supplier. Therefore, if there are several contracts between the saIne supplier and departrnent, \ve know that the saIne part Inus!; be involved in all of thcrn. This constraint is an FD, $DS \rightarrow P$.

Again we have redundancy and its associated problclns. We can address this situation by decornposing Contracts into two relations with attributes *CQSD* and *3DP.* Intuitively, the relation *3DP* records the part supplied to a depart-rllent by a supplier, and the relation *C:QSD* records additional infornlation about a contract. It is unlikely that we would arrive at such a design solely through ER rIlodeling, sinee it is hard to fOfrnulate an entity or relationship that corresponds naturally to *CQSD.*

19.7.3 Identifying Attributes of Entities

This exarIlple illustrates how a careful examination of FDs can lead to a better understanding of the entities and relationships underlying the relational tables; in particular, it shows that attributes can easily be associated with the 'wrong' entity set during ER design. The ER diagrarn in Figure 19.11 shows a relationship set called Works_In that is similar to the Works.ln relationship set of Chapter 2 but with an additional key constraint indicating that an employee can work in at rnost one departrIlent. (Observe the arrow connecting Employees to Works_In.)

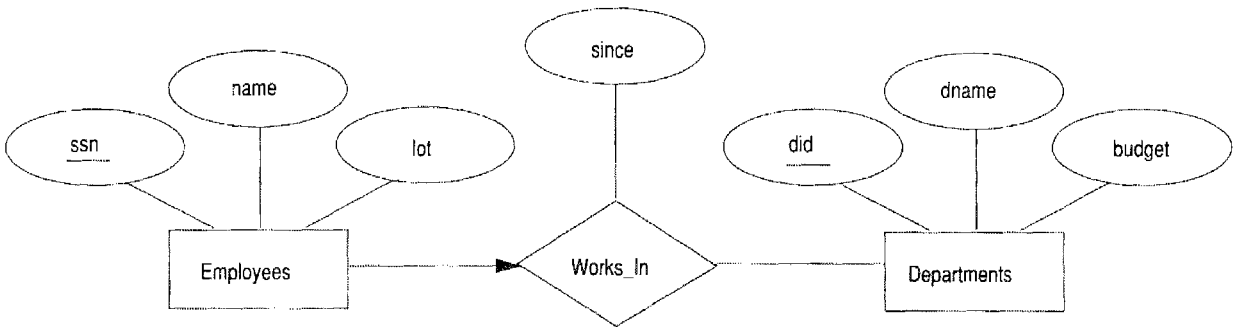


Figure 19.11. The Works._In Relationship Set

Using the key constraint, we can translate this ER diagrarn into two relations:

Workers(*ssn, narne, lot, d'id, since*)
Departrnents(*did, dname, budget*)

The entity set Ernployees and the relationship set Works_In are rnapped to a single relation, vVorkers. This translation is based on the second approach discussed in Section 2.4.1.

Now suppose elIlployees are assigned parking lots based on their departrnent, and that all enlployees in a given departrnent are assigned to the saIne lot. This constraint is not expressible with respect to the ER, diagrarIl of Figure 19.1l. It is another example of an FD: *did* → *lot.* The redundancy in this design can be elirninated by decornposing the Workers relation into two relations:

vVorkers2(*ssn, name, did, since)*
Dept_Lots(*did, lot)*

'rhe new design has lnuch to reconunend it. We can change the lots associated with a departlnent by updating a single tuple in the second relation (i.e., no update anornalies). We can associate a lot with a department even if it currently has no crnployees, without using *null* values (i.e., no deletion anornalies). We can add an eruployee to a department by inserting a tuple to the first relation even if there is no lot associated with the enlployee's departrnent (i.e., no insertion anornalies).

Exalnining the two relations Departrnents and Dept_Lots, which have the saIne key, we realize that a Departrnents tuple and a Dept_Lots tuple with the sarne key value describe the sarne entity. This observation is reflected in the ER cliagrarn shown in Figure 19.12.

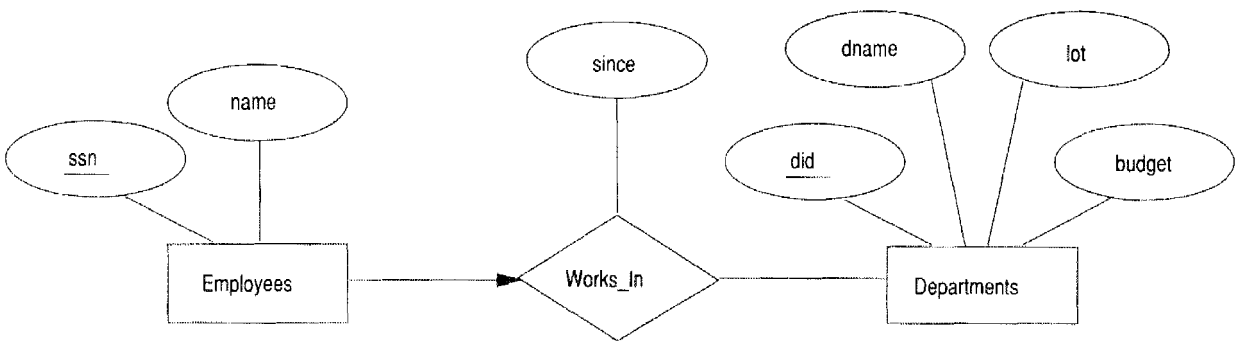


Figure 19.12 Refined\Norks_In Relationship Set

Translating this diagrarn into the relational rnodel would yield:

Workers2(*8871" narne, did, since)*
I)epartrnentsCdid, *dname, budget, lot)*

It SeCllS intuitive to associate lots with crnployees; on the other hand, the les reveal tllat ,in this exarnple lots are really associated with departrnents. The subjective process of ER modeling could Iniss this point. T'he rigorous process of norrnaliza,tion would not.

19.7.4 Identifying Entity Sets

Consider a variant of the Reserves scherna used in earlier chapters. Let Reserves contain attributes S, B, and D as before, indicating that sailor S has a reservation for boat B on day D. In addition, let there be an attribute C denoting the credit card to which the reservation is charged. We use this exarnple to illustrate how FD illfonnation can be used to refine an ER design. In particular, we discuss how FD inforluation can help decide whether a concept should be rnodeled as an entity or as an attribute.

Suppose every sailor uses a unique credit card for reservations. This constraint is expressed by the FD $S \rightarrow C$. This constraint indicates that, in relation Reserves, we store the credit card rnllnber for a sailor as often as we have reservations for that sailor, and we have redundancy and potential update anolnalies. A solution is to deconlpose Reserves into two relations with attributes SBD and SC. Intuitively, one holds inforrnation about reservations, and the other holds infonnation about credit cards.

It is instructive to think about an ER design that would lead to these relations. One approach is to introduce an entity set called Credit_Cards, with the sale attribute *cardno,* and a relationship set Has_Card associating Sailors and Credit_Cards. By noting that each credit card belongs to a single sailor, we can Inap Has_Card and Credit_Cards to a single relation with attributes SC. We would probably not rnodel credit card nUlnbers as entities if our Inain interest in card nurnbers is to indicate how a reservation is to be paid for; it suffices to use an attribute to rnodel card nUlnbers in this situation.

A second approach is to rnake *cardno* an attribute of Sailors. But this approach is not very natural--a sailor lllay have several cards, and we are not interested in all of theln. Our interest is in the one card that is used to pay for reservations, which is best lnodeled as an attribute of the relationship Reserves.

A helpful way to think about the design problern in this exarnple is that we first lnake *cardno* an attribute of H,eserves and then refine the resulting tables by taking into account the FD information. (Whether we refine the design by adding *cardno* to the table obtained froTll Sailors or by creating a new table with attributes SC is (1, separate issue.)

19.8 OTHER KINDS OF DEPENDENCIES

FI)s are probal)l.y the rn08t conunon and irnportant kind of constraint fr0 III the point of view of database design. However, there are several other kinds of dependencies. In particular, there is a well-developed theory for database

design llsing *multivalued dependencies* and *join dependencies.* By taking sueh dependencies into account, we can identify potential redundancy problenls that cannot be detected using FDs alone.

'rhis section illustrates the kinds of redundancy that can be detected using II1Ul-tivalued dependencies. Our Inain observation, however, is that sirnple guidelines (which can be checked using only FD reasoning) can tell us whether we even need to worry about complex constraints such as 111ultivalued and join dependencies. We also conunent on the role of *inclusion dependencies* in database design.

19.8.1 Multivalued Dependencies

Suppose that we have a relation with attributes *course, teacher,* and *book,* which we denote as *CTB.* The Ineaning of a tuple is that teacher *T* can teach course *C,* and book *B* is a reccnnmended text for the course. There are no FDs; the key is *CTB.* However, the recolnlnended texts for a course are independent of the instructor. The instance shown in Figure 19.13 illustrates this situation.

course	*teacher*	*book*
Physics101	Green	Mechanics
PhysicslOl	Green	Optics
PhysicslOl	Brown	Mechanics
Physics101	Brown	Optics
Math301	Green	Mechanics
Math301	Green	Vectors
Math301	Green	Geometry

Figure 19.13 BCNF R.elation with Redundancy That Is Revealed by MVDs

Note three points here:

- The relation sehcrna *CTB* is in BCNF; therefore we would not consider decolnposing it further if we looked only at the FDs that hold over *(JTB.*
- There is redundancy. rrhe fact that G·reen can teach Physics101 is recorded once per recommended text for the course. Sirnila.rly, the fact that Optics is a text for Physics101 is recorded once per potential teacher.
- T'he redundancy can be elirninated by decornposing *CTB* into CT and *CE.*

The redundaJ1cy in this exarnple is due to the constraint that the texts for a course are independent of tIle instructors, which cannot be expressed in tenns

of FDs. rrhis constraint is an example of a *multivalued dependency*, or MVD. Ideally, we should rnodel this situation using two binary relationship sets, Instructors with attributes CT and Text with attributes CB. Because these are two essentially independent relationships, rnodeling them with a single ternary relationship set with attributes CTE is inappropriate. (See Section 2.5.3 for a further discussion of ternary versus binary relationships.) Given the subjectivity of ER design, ho\vever, we rnight create a ternary relationship. A careful analysis of the fvIVD infonnation would then reveal the problern.

Let R be a relation scheIna and let X and Y be subsets of the attributes of R. Intuitively, the multivalued dependency $X \rightarrow\rightarrow$ Y'is said to hold over R if, in every legal instance r of R, each X value is associated with a set of Yvalues and this set is independent of the values in the other attributes.

ForInally, if the MVD $X \rightarrow\rightarrow Y$ holds over R and $Z = R - XY$, the following lllUSt be true for every legal instance r of R:

> If tl E r, $t2$ E rand $t1.X = t2.X$, then there must be sorne $t3$ E r such that $t1'XY = t3.XY$ and $t2 \cdot Z = t3'Z$,

Figure 19.14 illustrates this definition. If we are given the first two tuples and told that the MVD $X \rightarrow\rightarrow Y$ holds over this relation, we can infer that the relation instance must also contain the third tuple. Indeed, by interchanging the roles of the first two tuples—treating the first tuple as $t2$ and the second tuple as t_1—we can deduce that the tuple $t4$ must also be in the relation instance.

X	Y	Z	
a	b_1	C1	tuple $t1$
a	b_2	C2	tuple $t2$
a	b_1	C2	tuple t_3
a	b_2	C1	tuple $t4$

Figure 19.14 Illustration of MVD Definition

This table suggests another way to think about lVIVDs: If $X \rightarrow\rightarrow Y$ holds over R, then $\pi_{YZ}(\sigma_{X=x}(R)) = \pi_Y(\sigma_{X=x}(R)) \times \pi_Z(\sigma_{X=x}(R))$ in every legal instance of R, for any value x that appears in the X colurnn of R. In other words, consider groups of tuples inR with the sarne X-value. In each such group consider the projection onto the attributes YZ. This projection HUlst be equal to the cross-product of the projectiolls onto Yand Z. That is, for a given X-value, the Y-values and Z-values are independent. (Froln this definition it is easy to see that $X \rightarrow\rightarrow$ y'l11Ust hold wherlever $X \rightarrow Y$ holds. If the FI) $X \rightarrow$

Y holds, there is exactly one Y-value for a given X-value, and the conditions in the MVD definition hold trivially. The converse does not hold, as Figure 19.14 illustrates.)

Returning to our CTB exalllple, the constraint that course texts are independent of instructors can be expressed as $C \twoheadrightarrow T$. In terlllS of the definition of MVDs, this constraint can be read as follo\vs:

> If (there is a tuple showing that) C is taught by teacher T,
> and (there is a tuple showing that) G has book B as text,
> then (there is a tuple showing that) G is taught by T and has text B.

Given a set of FDs and MVDs, in general, we can infer that several additional FDs and MVDs hold. A sound and complete set of inference rules consists of the three ArIllstrong AxioIllS plus five additional rules. Three of the additional rules involve only MVDs:

- MVD Complementation: If $X \twoheadrightarrow Y$, then $X \twoheadrightarrow R - XY$.
- MVD .Augmentation: If $X \twoheadrightarrow$ Yand $W \supseteq Z$, then $WX \twoheadrightarrow YZ$.
- MVD Transitivity: If $X \twoheadrightarrow$ Yand $Y \twoheadrightarrow Z$, then $X \twoheadrightarrow (Z - Y)$.

As an exanlple of the use of these rules, since we have $C \twoheadrightarrow T$ over GTB, MVD complelnentation allows us to infer that $C \twoheadrightarrow OTB - CT$ as well, that IS, $C \twoheadrightarrow B$. The remaining two rules relate FDs and MVDs:

- Replication: If $X \rightarrow$ Y, then $X \twoheadrightarrow Y$.
- Coalescence: If $X \twoheadrightarrow$ Yand there is a W such that $W \cap Y$- is elnpty, $W \rightarrow Z$, and $Y \supseteq Z$, then $X \rightarrow Z$.

()bserve that replication states that every FD is also an MVD.

19.8.2 Fourth Normal Form

Fourth Horrnal fonn is a direct generalization of BeNF. Let R be a relation scherna, X and Y be nonernpty subsets of the attributes of R, and F be a set of dependencies that includes both FDs and MVDs. R is said to be in fourth normal form (4NF), if, for every I\1.VI) $X \twoheadrightarrow Y$ that holds over R, one of the following staternents is true:

- $Y \subseteq X$ or $XY = R$, or
- X is a superkey.

In reading this definition, it is irnportant to understand that the deflnition of a *key* has not changed·····-thkey rnust uniquely deterrnine all attributes through FDs alone. $X \rightarrow\rightarrow$ Y'is a trivial MVD if $Y \subset X \subseteq R$ or $XY = R$; such MVDs always hold.

The relation *CTB* is not in 4NF because $C \rightarrow\rightarrow T$ is a nontrivial MVD and *C* is not a key. We can elirninate the resulting redundancy by deconlposing *CTB* into *CT* and *CB;* each of these relations is then in 4NF.

To use MVD inforrnation fully, we nUlst understand the theory of MVDs. However, the following result due to Date and Fagin identifies conditions-detected using only FD information!—under which we can safely ignore MVD inforrnation. That is, using MVD information in addition to the FD infornlation will not reveal any redundancy. Therefore, if these conditions hold, we do not even need to identify all MVDs.

> If a relation schema is in BCNF, and at least one of its keys consists of a single attribute, it is also in 4NF.

An in1.portant assl.unption is inlplicit in any application of the preceding result: *The set of FDs identified thus far is 'indeed the set of all FDs that hold over the relation.* This assulllption is important because the result relies on the relation being in BCNF, which in turn depends on the set of FDs that hold over the relation.

We illustrate this point using an exalnple. Consider a relation scherna *ABCD* and suppose that the FD $A \rightarrow BCD$ and the MVD $B \rightarrow\rightarrow C$ are given. Considering only these dependencies, this relation schema appears to be a counterexalnple to the result. The relation has a sirnple key, appears to be in BCNF, and yet is not in 4NF because $B \rightarrow\rightarrow C$ causes a violation of the 4NF conditions. Let us take a closer look.

B	*C*	*A*	*D*	
b	C1	0:1	d_1	tuple t_1
b	C2	([,2	d_2	tuple t_2
b	C1	([,2	*d2*	tuple t_3

Figure 19.15 Three Tuples [rorn a Legal Instance of *ABCD*

Figure 19.15 8ho\v8 three tuples f1'orn an instance of *ABCD* that satisfies the given MVD $B \rightarrow\rightarrow C$. Frolu the definition of an MVD, given tuples tl and t_2, it follows that tuple t_3 Inust also be included in the instaJ1ce. Consider tuples t_2 and t_3. FrOIJI the givenFD $A \rightarrow BCD$ and the fact that these tuples have the

same A-value, we can deduce that c1 = c2. Therefore, we see that the FD $B \rightarrow C$ rnust hold over $ABCD$ whenever the FD $A \rightarrow BCD$ and the MVD $B \rightarrow\rightarrow C$ hold. If $B \rightarrow C$ holds, the relation $ABeD$ is not in BeNF (unless additional FDs lllake B a key)!

Thus, the apparent counterexalnple is really not a counterexalllple·····-rather, it illustrates the iInportance of correctly identifying all FDs that hold over a relation. In this exarnple, $A \rightarrow BCI)$ is not the only FD; the FD $B \rightarrow C$ also holds but ·was not identified initially. Given a set of FDs and IvIVI)s, the inference rules can be used to infer additional FDs (and l\1VDs); to apply the Date-Fagin result without first using the l\1VD inference rules, we IUUSt be certain that we have identified all the FDs.

In summary, the Date-Fagin result offers a convenient way to check that a relation is in 4NF (without reasoning about l\1VDs) if we are confident that we have identified all FDs. At this point, the reader is invited to go over the examples we have discussed in this chapter and see if there is a relation that is not in 4NF.

19.8.3 Join Dependencies

A join dependency is a further generalization of MVDs. A join dependency (JD) $\bowtie \{R_1, \ldots, R_n\}$ is said to hold over a relation R if $R_1, \ldots, R_n$ is a lossless-join decolnposition of R.

An MVD $X \rightarrow\rightarrow$ Yover a relation R can be expressed as the join dependency $\bowtie$ {XV, X(R,--Y)}..As an example, in the GTB relation, the MVD C $\rightarrow\rightarrow T$ can be expressed as the join dependency $\bowtie$ {Crr, CB}.

U·nlike FDs and l'v1VDs, there is no set of sound and cornplete inference rules for JDs.

19.8.4 Fifth Normal Form

A relation schcrna R is said to be in fifth normal form (5NF) if, for every .II) $\bowtie \{.R_1, \bullet\bullet\bullet, R_n\}$ that holds over R, one of the follo"ving statcrnents is true:

- $R_i = R$, for scnne i, or
- The $.lD$ is irnplied by the set of those FDs over R in ·which the left side is a key for R.

The second condition deserves S(Hne explanation, since we have not presented inference rules for FDs and .00Ds taken together. Intuitively, we rnust be able to sho\v that the decolnposition of R into $\{R_1, \ldots, R_n\}$ is lossless-join whenever the key dependencies (FDs in which the left side is a key for R) hold. JI) $\bowtie \{R_1, \ldots, R_n\}$ is a trivial JD if $R_i = R$ for SaIne i; such a .JD always holds.

The following result, also due to Date and Fagin, identifies conditions--- again, detected llsing only FD inforlnation---under -which we can safely ignore JD inforlnation:

> If a relation schenla is in 3NF and each of its keys consists of a single attribute, it is also in 5NF.

The conditions identified in this result are sufficient for a relation to be in 5NF but not necessary. rrhe result can be very useful in practice because it allows us to conclude that a relation is in 5NF *'Without ever 'identifying the MVDs and JDs that 'may hold oveT the relation.*

19.8.5 Inclusion Dependencies

lVIVDs and JDs can be used to guide database design, as we have seen, although they are less COllUllon than FDs and harder to recognize and reason about. In contrast, inclusion dependencies are very intuitive and quite cornrnon. IIowever, they typically have little influence on database design (beyond the ER design stage).

Infonnally, an inclusion dependency is a statement of the fOlTH that S0111e cohunns of a relation are contained in other cohunns (usually of a second relation). A foreign key constraint is an example of an inclusion dependency; the referring colurnn(s) in one relation rnust be contain.ecl in the prirnary key cohnnn(s) of the referenced relation. As another exarnple, if!? and S are two relations obtained 1)y translating two entity sets that every R entity is also an S erltity, we would have an inclusion dependency; projecting R on its key attributes yields a relation contained in the relation obtained by projecting S on its key attributes.

The rnain point to bear in rnind is that we should not split groups of attributes that participate in an inclusion dependency. For exarnple, if we have an inclusion dependency $AB \subseteq Of)$, \vhile decornposing the relation scherna containing AB, we should ensure that at least one of the schemas obtained in the deccnnposition contains botJ1 A and B. Otherwise, we cannot check the inclusion clependency $AB \subseteq CD$ without reconstructing the relation containing $A13$.

Ivlost inclusion dependencies in practice are *key-based*, that is, involve only keys. Foreign key constraints are a good exalnple of key-based inclusion dependencies. An ER diagram that involves ISA hierarchies (see Section 2.4.4) also leads to key-based inclusion dependencies. If all inclusion dependencies are key-based, we rarely have to worry about splitting attribute gTOUps that participate in inclusion dependencies, since decornpositions usually do not split the primary key. N'ote, however, that going fn:)111 3NF to BCNF always involves splitting SOlne key (ideally not the prirnary key!), since the dependency guiding the split is of the fornl $X \rightarrow A$ where A is part of a key.

19.9 CASE STUDY: THE INTERN'ET SHOP

R,ecall froIn Section 3.8 that DBDudes settled on the following scherna:

Books(*isbn:* CHAR(10), *title:* CHAR(8) , *author:* CHAR(80) ,
qty_in_stock: INTEGER, *price:* REAL, *year_published:* INTEGER)
Custolllers(*cid:* INTEGER, *cnaTne:* CHAR(80) , *address:* CHAR(200))
Orders(*orde.rnum,:* *INTEGER*, *isbn:* CHAR(.10), *cid:* INTEGER,
cardnu'm: CHAR(16), *qty:* INTEGER, *ordeT_date:* DATE, *ship_date:* DATE)

DBDudes analyzes the set of relations for possible redundancy. The Books relation has only one key, *(isbn),* and no other functional dependencies hold over the table. Thus, Books is in BCNF. The Custorners relation also has only one key, (cid), and no other functional depedencies hold over the table. T'hus, Custorners is also in BCNF.

DBI)udes has already identified the pair $\langle ordernum,\ isbn \rangle$ as the key for the Orders table. In addition, since each order is placed by one custorner on one specific date with one specific credit card nurnber, the following three functional dependencies hold:

$ordernum \rightarrow cid$, $ordernum \rightarrow order\_date$, and $ordernum \rightarrow cardnum$

The experts at DBDudes conclude that Orders is not even in 3NF. (Can you see why?) They decide to clecornpose ()rders into the following two relations:

Orders(*ordernum*, *cid*, *order_date*, *cardnum*, a.nd
()rderlists(*ordernum*, *isbn*, *qty*, *ship_date*)

The resulting t\vo relations, ()rders and ()rderlists, are both in BCNF', and the decornposition is lossless-join since *ordernum* is a key for (the new) ()rders. The reader is invited to check that this decolnposition is also dependency-preserving. For cornpleteness, we give th,e SQL DIJL for the ()rders and Orderlists relations below:

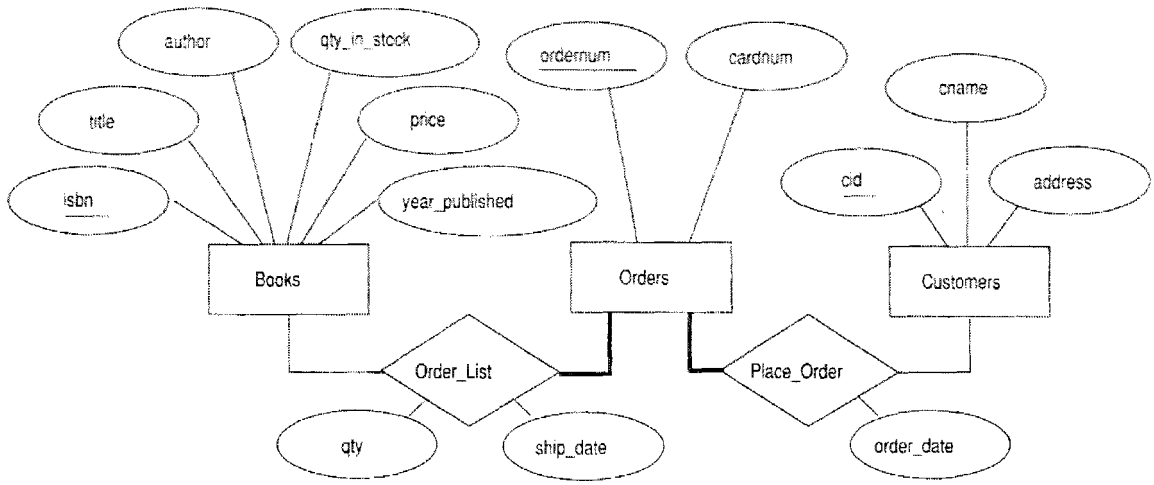


Figure 19.16 ER Diagram Reflecting the Final Design

```
CREATE TABLE Orders ( ordernurn     INTEGER,
                      cid           INTEGER,
                      order_date    DATE,
                      cardnum       CHAR(16),
                      PRIMARY KEY   (ordernllllll),
                      FOREIGN KEY   (cid) REFERENCES  Custolllers )

CREATE TABLE Orderlists (ordernurll   INTEGER,
                         isbn         CHAR(10),
                         qty          INTEGER,
                         ship_date    DATE,
                         PRIMARY KEY  (ordernurn, isbn),
                         FOREIGN KEY  (isbn) REFERENCES  Books)
```

F'igure 19.16 shows an updated ER diagrarn that reflects the new design. Note that DBDudes could have arrived inunedia,tely at this diagrarn if they had made ()rders an entity set instead of a relationship set right at the beginning. But at that tilne they did not understand the requirernents cornpletely, and it seeTHed natural to rnodel Orders as a relationship set. This iterative refinernent process is typical of real-life da,tabase design processes. As DBI)udes has learned over tirne, it is rare to achieve an initial design that is not changed as a project progresses.

T'he DBI)udes team celebrates the successful cornpletion of logical database design and scherna refinelnent by opening a bottle of charnpagne and charging it to B&:N. After recovering frorn the celebration, they lIlove on to the physical design phase.

19.10 REVIEW QUESTIONS

Answers to the review' questions can be found in the listed sections.

- Illustrate redundancy and the problerns that it ca.n cause. Give examples of *insert*, *delete*, and *update* anoInalies. Can *null* values help address these problerlls? Are they a cOlnplete solution? (Section 19.1.1)
- What is a *decoTnpositio'n* and how does it address redundancy? What problerlls Inay be caused by the use of decolupositions? (Sections 19.1.2 and 19.1.3)
- Define *functional dependencies.* How are *prtmary keys* related to FDs? (Section 19.2)
- When is an PD *j implied by* a set *F* of FDs? Define *Armstrong's Axioms*, and explain the statement that "they are a sound and cornplete set of rules for FD inference." (Section 19.3)
- What is the *dependency closure F+* of a set *F* of FDs? What is the *attribute closure X+* of a set of attributes *X* with respect to a set of FDs *F?* (Section 19.3)
- Define INF, 2NF, 3NF, and BCNF. What is the nlotivation for putting a relation in BCNF? What is the motivation for 3NF? (Section 19.4)
- When is the decomposition of a relation schenla *R* into two relation schemas *X* and *Y* said to be a *lossless-join* decomposition? Why is this property so irnportant? Give a necessary and sufficient condition to test whether a decc)1nposition is lossless-join. (Section 19.5.1)
- When is a decornposition said to be *depc'ndency-preserving?* Why is this property useful? (Section 19.5.2)
- Describe how we can obtain **a** lossless-join decornposition of a relation into BCNF. Give an exanlple to show that there rnay not be a dependency-preserving decornposition into BCNF. Illustrate how a given relation could be decornposed in different ways to arrive at several alternative decornpositions, and discuss the irnplications for database design. (Section 19.6.1)
- Give an example that illustrates how a collection of relations ill BCNF could have redundancy even though each relation, by itself, is free fronl redundancy. (Section 19.6.1)
- What is a. *Tninirnal cover* for a set of FDs? Describe an algorithrn for cornputing the minimal cover of B set of FI)s, and illustrate it with an exarnple. (Section 19.6.2)

- Describe how the algorithl11 for lossless-join decolnposition into BCNF can be adapted to obtain a lossless-join, dependency-preserving decornposition into 3NF. Describe the alternative *synthesis* approach to obtaining such a decorllposition into 3NF. Illustrate both approaches using an exarnple. (Section 19.6.2)
- Discuss how scherna refinernent through dependency analysis and norrnalization can iInprove schemas obtained through ER design. (Section 19.7)
- Define *multivalued dependencies, join dependencies,* and *inclusion dependencies.* Discuss the use of such dependencies for database design. Define 4NF and 5NF, and explain how they prevent certain kinds of redundancy that BCNF does not eliminate. Describe tests for 4NF and 5NF that use only FDs. What key assumption is involved in these tests? (Section 19.8)

EXERCISES

Exercise 19.1 Briefly answer the following questions:

1. Define the term *functional dependency.*
2. Why are some functional dependencies called *trivial*?
3. Give a set. of FDs for the relation schema $R(A,B,C,Dj$ with prilnary key AB under which R is in 1NF but not in 2NF.
4. Give a set of FDs for the relation schelna $R(A,B,C,Dj$ with prilnary key AB under which R is in 2NF but not in 3NF.
5. Consider the relation schelna $R(A,B,OJ$, which has the FD $B \rightarrow$ C. If A is a candidate key for R, is it possible for R to be in BCNF? If so, under what conditions? If not, explain why not.
6. Suppose we have a relation schema $R(A,B,OJ$ representing a relationship between two entity sets with keys A and 13, respectively, and suppose that R has (aIIlong others) the FDs $A \twoheadrightarrow$ Band $13 \rightarrow A$. Explain what such a pair of dependencies means (i.e., what they irnply about the relationship that the relation nlOdels).

Exercise 19.2 Consider a relation R with five attributes $ABCDE$. You are given the follc)\ving dependencies: $A \rightarrow B$, $Be \rightarrow E$, and $ED \rightarrow A$.

1. List all keys for R.
2. Is R in 3NF?
3. Is R in BCNF?

Exercise 19.3 Consider the relation shown in Figure 19.17.

1. List all the functional dependencies that this relation instance satisfies.
2. Assume that the value of attribute Z of the last record in the relation is changed frorH z_3 to z_2. Now list all the functional dependencies that this relation instance satisfies.

Exercise 19.4 Assurne that you are given a relation with attributes $ABCD$.

X	Y	Z
x1	y_1	z1
x1	Y1	z2
x2	Y1	z1
x_2	Y1	z_3

Figure 19.17 Relation for Exercise 19.3.

1. Asslune that no record has NULL values. \Nrite an SQL query that checks whether the functional dependency $A \rightarrow B$ holds.
2. Assulne again that no record has NULL values. Write an SQL assertion that enforces the functional dependency $A \rightarrow B$.
3. Let us now aSSUlne that records could have NULL values. Repeat the previous two questions under this assurnption.

Exercise 19.5 Consider the following collection of relations and dependencies. Assume that each relation is obtained through decomposition from a relation with attributes *ABCDEFGHI* and that all the known dependencies over relation *ABCDEFGHI* are listed for each question. (The questions are independent of each other, obviously, since the given dependencies over *ABCDEFGHI* are different.) For each (sub)relation: (a) State the strongest nonnal fonn that the relation is in. (b) If it is not in BCNF, decompose it into a collection of BCNF relations.

1. *R1(A. C,B.D,E), A → 13, C → D*
2. *R2(A,B,F), AC → B, B → F*
3. *R3(A,D), D → G, G → H*
4. *R4(D, C,H, G), A → I, I → A*
5. *R5(A.I,C.B)*

Exercise 19.6 Suppose that we have the following three tuples in a legal instance of a relation schema *S* with three attributes *ABC* (listed in order): (1,2,:3), (4,2,3), and (5,3,3).

1. \Vhich of the following dependencies can you infer does *not* hold over scherna *S*?

   (a) $A \rightarrow 13$, (b) $Be \rightarrow A$, (c) $13 \rightarrow C$

2. Can you identify allY dependencies that hold over *S*?

Exercise 19.7 Suppose you are given a relation R with four attributes $ABCD$. For each of the following sets of FDs, assurning those are the only dependencies that hold for R, do the following: (a) Identify the candidate key(s) for *R*. (b) Identify the best Honnal forBl that R satisfies (1NF, 2NF, 3NF, or BeNF). (c) If R is not in BCNF, decOlnpose it into a set of BCNF relations that preserve the dependencies.

1. $C \rightarrow D$, $C \rightarrow A$. 13 C
2. $B \rightarrow C$. $D \rightarrow A$
3. $ABC \rightarrow D$, $D \rightarrow A$
4. $A \rightarrow B$. $BC \rightarrow D$. $A \rightarrow C$
5. $A13 \rightarrow C$, $AB \rightarrow D$. $C \rightarrow A$, $D \rightarrow 13$

Exercise 19.8 Consider the attribute set $R = ABCDEGH$ and the FD set $F = \{AB \rightarrow C, AC \rightarrow B, AD \rightarrow E, B \rightarrow D, BE \rightarrow A, B \rightarrow G\}$.

1. For each of the following attribute sets, do the following: Cornpute the set of dependencies that hold over the set and write down a rninirnal cover. (ii) Narne the strongest nonnal [onn that is not violated by the relation containing these attributes. (iii) DeCOlnpose it into a collection of BCNF relations if it is ıH)t in BeNF'.

   (a) *ABC,* (b) *ABCD,* (c) *ABCEG,* (d) *DC:BGII,* (e) *ACEH*

2. Which of the following decOIllpositions of $R = ABCDEG$, with the saIne set of dependencies F, is (a) dependency-preserving? (b) lossless-join?

   (a) $\{AB, BC, ABDE. EG\}$

   (b) $\{ABC, ACDE, ADG\}$

Exercise **19.9** Let R be decOIllposed into $R_1, R_2, \ldots, R_n$. Let F be a set of FDs on R.

1. Define what it rlleans for F to *be pre8erved* in the set of decOlllposed relations.
2. Describe a polynomial-tirne algorithm to test dependency-preservation.
3. Projecting the FDs stated over a set of attributes X onto a subset of attributes Y requires that we consider the closure of the FDs. Give an exarnple where considering the closure is irnportant in testing dependency-preservation, that is, considering just the given FDs gives incorrect results.

Exercise **19.10** Suppose you are given a relation $R(A,B,C,D)$. For each of the following sets of FDs, assuming they are the only dependencies that hold for R, do the following: (a) Identify the candidate key(s) for R. (b) State whether or not the proposed decOlnposition of R into smaller relations is a good decolllposition and briefly explain why or why not.

1. $B \rightarrow C$, $D \rightarrow A$; decornpose into BC and AD.
2. $AB \rightarrow C$, $C \rightarrow A$, $C \rightarrow D$; decompose into ACD and BE.
3. $A \rightarrow BC$, $C \rightarrow AD$; decornpose into ABC and AD.
4. $A \rightarrow B$, B C, C $\rightarrow D$; decornpose into AB and ACD.
5. $A \rightarrow B$, $B \rightarrow C$, $C \rightarrow D$; decOInpose into AB, AD and CD.

Exercise 19.11 Consider a relation R that has three attributes ABC. It is decornposed into relations R_1 with attributes AB and R_2 with attributes BC.

1. State the definition of a lossless-join decOlnposition with respect to this exarnple. Answer this question concisely by writing a relational algebra equation involving R, R_1, and $R2$.
2. Suppose that $B \twoheadrightarrow C$. Is the decorHposition of R into R_1 and $R2$ lossless-join? Reconcile your answer with the observation that neither of the FDs HI nR2 $\rightarrow$ R I nor R I n $R_2 \rightarrow R_2$ hold, in light of the simple test offering a necessary and sufficient condition for lossless-join decmnposition into two relations in Section 15.6.1.
3. If you are given the following justa.nees of R_1 a,nd R_2, what can you say about the instance of R from which these were obtained? Answer this question by listing tuples that are definitely ill R and tuples that are possibly in R.

   Instance of R_1 = {(5,1), (6,1)}
   Instance of R_2 = {(1,8), (1,9)}

   Can you say that attribute B definitely *is* or *is not* a key for R?

Exercise 19.12 Suppose that we have the following four tuples in a relation S with three attributes *ABC:* (1,2,3), (4,2,3), (5,3,3), (5,3,4). Which of the following functional ($\rightarrow$) and rIlultivalued ($\rightarrow\!\!\!\rightarrow$ dependencies can you infer does *not* hold over relation S?

1. $A \rightarrow 13$
2. $A \rightarrow\!\!\!\rightarrow$ B
3. $Be \rightarrow A$
4. $BG \rightarrow\!\!\!\rightarrow A$
5. 8 $\rightarrow C$
6. $B \rightarrow\!\!\!\rightarrow$ C

Exercise 19.13 Consider a relation R with five attributes $ABCDE$.

1. For each of the following instances of R, state whether it violates (a) theFD $Be \rightarrow D$ and (b) the MVD $Be \rightarrow\!\!\!\rightarrow D$:
   (a) { } (i.e., mnpty relation)
   (b) {(0,,2,3,4,5), (2,a,3,5,5)}
   (c) {(o,,2,3,4,5), (2,0,,3,5,5), (o,,2,3,4,6)}
   (d) {(a,2,3,4,5), (2,0,,3,4,5), (o,,2,3,6,5)}
   (e) {(0,,2,3,4,5), (2,0,,3,7,5), (a,2,3,4,6)}
   (f) {(o,,2,3,4,5), (2,0,,3,4,5), (0,,2,3,6,5), (o,,2,3,6,6)}
   (g) {(a,2,3,4,5), (0,,2,:3,6,5), (0,,2,3,6,6), (o,,2,3,4,6)}
2. If each instance for R listed above is legal, what can you say about the FD $A \rightarrow B$?

Exercise 19.14 JDs are lllotivated by the fact that sornetilnes a relation that cannot be decoruposed into two sinaller relations in a lossless-join rnanner can be so deC0111pOsed into three or rnore relations. An example is a relation with attributes *supplier*, *part*, and *project*, denoted *SPJ*, with no FDs or MVDs. The JD $\bowtie$ {*SP*, *PJ*, *JS*} holds.

Frorn the JD, the set of relation scheines *SP*, *PJ*, and *JS* is a IORsless-join decornposition of *SPJ*. Construct an instance of *HPJ* to illustrate that no two of these schernes suffice.

Exercise 19.15 Answer the following questions

1. Prove that the algorithrn shown in Figure 19.4 correctly cornputes the attribute closure of the input attribute set X.
2. Describe a linear-tirne (in the size of the set of FI)s, where the size of each FD is the nurnber of attributes involved) algoritlun for finding the attribute closure of a set of attributes with respect to a set of FDs. Prove that your algoritlun correctly COInputes the attribute closure of the input attribute set.

Exercise 19.16 Let us say that an 'Fl) $X \rightarrow Y$ is *simple* if Y is a single attribute.

1. Replace the FD $AB \rightarrow CD$ lJy the srnallest equivalent collection of sirnple FDs.
2. Prove that everyFD $X \rightarrow Y$ in a set of FDs F can be replaced by a set of sirnple F'Ds such that F^+ is equal to the closure of the new set of FDs.

Exercise 19.17 Prove that Arrnstrong's Axioms are sound and complete for FD inference. That is, show that repeated application of these axioms on a set F ofF'Ds produces exactly the dependencies in $P+$.

Exercise 19.18 Consider a relation R with attributes $ABCDE$. Let the following FDs be given: $A \rightarrow BC$, $Be \rightarrow E$, and $E \rightarrow DA$. Siluilarly, let S be a relation with attributes $ABCDE$ and let the follo\ving FDs be given: $A \rightarrow BC$, $B \rightarrow E$, and $E \rightarrow DA$. (Only the second dependency differs frolll those that hold over R.) You do not know whether or which other (join) dependencies hold.

1. Is R in BCNF?
2. Is R in 4NF?
3. Is R in 5NF?
4. Is Sin BeNF?
5. Is Sin 4NF?
6. Is Sin 5NF'?

Exercise 19.19 Let R be a relation schelna with a set F of FDs. Prove that the decOIll-position of R into HI and $R2$ is lossless-join if and only if $p+$ contains HI $\cap R_2 \rightarrow R_1$ or $R_1 \cap R2 \rightarrow R_2$.

Exercise 19.20 Consider a scheme R with FDs F that is decOlnposed into schelnes with attributes X and Y. Show that this is dependency-preserving if $F' \subseteq (F_X \cup py)+$.

Exercise 19.21 Prove that the optiInizatioll of the algorithrn for lossless··-join, dependency-preserving decornposition into 3NF relations (Section 19.6.2) is correct.

Exercise 19.22 Prove that the 3NF synthesis algoritlull produces a lossless-join decOlnposition of the relation containing all the original attributes.

Exercise 19.23 Prove that an MVD $X \twoheadrightarrow Y$ over a relation R can be expressed as the join dependency $\bowtie \{XY, X(R - Y)\}$.

Exercise 19.24 Prove that, if R has only one key, it is in BCNF if and only if it is in 3NF.

Exercise 19.25 Prove that, if R is in 3NF and every key is shnple, then R is in HeNF.

Exercise 19.26 Prove these staternents:

1. If a relation scherne is in BCNF and at least one of its keys consists of a single attrilmte, it is also in 4NF.
2. If a relation scherne is in 3NF and each key has a single attribute, it is also in 5NF'.

Exercise 19.27 Give an algorithrn for testing whether a relation scheme is in BCNF. The :.llgorithrn should l)e polynorniaJ in the size of the set of given FDs. (The *size* is the surn over all FI)s of the nurnber of attributes that appear in theFJ).) Is there a polyuOlnial algorithrn for testing whether a relation scheme is in 3NF?

Exercise 19.28 Give an algorithm for testing whether a relation scherne is in BCNF. The algorithm should be polynomial in the size of the set of given FI)s. (The 'size' is the SUln over all FI)s of the nurnber of attributes that appear in theFD.) Is there a polynomial algorithrn for testing whether a relation scheme is in 3NF?

Exercise 19.29 1)rove that the algorithm for decomposing a relation scherna with a set of FI)s into a collection of BCNS relation schemas as describerl in Section 19.6.1 is correct (i.e., it produces a collection of BCNF relations, and is lossless-join) and terrninates.

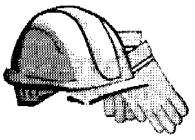
BIBLIOGRAPHIC NOTES

Textbook presentations of dependency theory and its use in database design include [3, 45, 501) 509, 747]. Good survey articles on the topic include [755, 415].

FDs were introduced in [187], along with the concept of 3NF, and aximlls for inferring FDs were presented in [38]. BCNF was introduced in [188]. The concept of a legal relation instance and dependency satisfaction are studied fonnaUy in [328].FDs were generalized to scrnantic data Illodels in [768].

Finding a key is shown to be NP-COlnplete in [497]. Lossless-join decOlnpositions were studied in [28, 502, 627]. Dependency-preserving decorIlpositions were studied in [74]. [81] introduced rninirnal covers. DecOlnposition into 3NF is studied by 1.81, 98] and decOlnposition into BCNF is addressed in [742]. [412] shows that testing whether a relation is in 3NF is NP-cOlnplete. [253] introduced 4NF and discussed decoillposition into 4NF. Fagin introduced other nonnal forIlls in [254] (project-join nonnal fonn) and [255] (doHluin-key nonnal forIll). **In** contrast to the extensive study of vertical decOlnpositions, there has been relatively little formal investigation of horizontal decompositions. [209] investigates horizontal decoillpositiolls.

IvIVDs were discovered independently by Delobel [211],Fagin [253], and Zaniolo [789]. AxiOlllS for FDs and MVDs were presented in [73}. [593] shows that there is no axiornatization for JDs, although [662] provides an axioHlatization for a more general class of dependencies. The sufficient conditions for 4NF and 5NF in tenns of FDs that were discussed in Section 19.8 are from [205]. An approach to database design that uses dependency inforrnation to construct sarnple relation instances is described in [508, 509].



20

PHYSICAL DATABASE DESIGN AND TUNING

- What is physical database design?
- What is a query workload?
- How do we choose indexes? What tools are available?
- What is co-clustering and how is it used?
- What are the choices in tuning a database?
- How do we tune queries and view?
- What is the impact of concurrency on perforrnance?
- How can we reduce lock contention and hotspots?
- What are popular database benchnlarks and how are they used?
- Key concepts: Physical database design, database tuning, workload, co-clustering, index tuning, tuning wizard, index configuration, hot spot, lock contention, database benchmark, transactions per second

Advice to a client who cornplained al)out rain leaking through the roof onto the dining table: "Move the table."

—Architect Frank Lloyd Wright

The perfonnance of a DBMS on cornrnonly asked queries and typical update operations is the ultirnate Ineasure of a database desigIl. A I}BA can irnprove perforrnance by identifying perforrnance bottlenecks and adjusting sorne DBIVlS pararneters (e.g., the size of the buffer pool or the frequency of checkpointing) or adding hardware to elirninate such bottlenecks. rIhe first step in achieving

good perforlnancc, however, is to Inake good database design choices, which is the focus of this chapter.

After we design the *conceptual* and *external* schernas, that is, create a collection of relations and views along 'with a set of integrity constraints, we .Il1ust address pe1'forlna11c8 goals through physical database design, in which we design the *physical* sche111a. As user requirernents evolve, it is usually necessary to tune, or adjust, all aspects of a database design for good perforrnance.

This chapter is organized as follows. We give an overview of physical database design and tuning in Section 20.1. The 1nost irnportant physical design decisions concern the choice of indexes. We present guidelines for deciding which indexes to create in Section 20.2. These guidelines are illustrated through several exalnples and developed further in Sections 20.3. In Section 20.4, we look closely at the irnportant issue of clustering; we discuss how to choose clustered indexes and whether to store tuples fron1 different relations near each other (an option supported by sorne DBMSs). In Section 20.5, we emphasize how well-chosen indexes can enable some queries to be answered without ever looking at the actual data records. Section 20.6 discusses tools that can help the DBA to autornatically select indexes.

In Section 20.7, we survey the lnain issues of database tuning. In addition to tuning indexes, we 111ay have to tune the conceptual schema as well as frequently used query and view definitions. We discuss how to refine the conceptual schelna in Section 20.8 and how to refine queries and view definitions in Section 20.9. We briefly discuss the perforrnance irnpact of concurrent access in Section 20.10. We illustrate tuning on our Internet shop exarnple in Section 20.11. We conclude the chapter with a short discussion of DBMS benchrnarks in Section 20.12; benchrnarks help evaluate the perfOI'lnanCe of alternative DBl\1S products.

20.1 INTRODUCTION TO PHYSICAL DATABASE DESIGN

Like all other aspects of database design, physical design rnust be guided by the nature of the data a,nd its intended use. In particular, it is irnportant tonnderstand the typical workload that the database lI1Ust support; the workload consists of a rnix of queries and updates. Users also have certain requirements about how fast certain queries or updates 11111st run or how rnan.y tran.sactions rnust be processed per second. The \vorkload description and users' performance reqllirernents are the basis on \vhich a nurnber of decisions have to be rnade during pl1ysical database design.

Identifying Perfornlance Bottlenecks: All commercial systems provide a suite of tools for rnonitoring a wide range of systeIll parameters. These tools, used properly, can help identify perforlnance bottlenecks and suggest aspects of the database design and application code that need to be tuned for perforlnance. For example, we can ask the DBMS to rnonitor the execution of the database for a certain period of tinle and report on the nurnber of clustered scans, open cursors, lock requests, checkpoints, buffer scans, average wait titne for locks, and many such statistics that give detailed insight into a *snapshot* of the live system. In Oracle, a report containing this inforlnation can be generated by running a script called UTLBSTAT. SQL to initiate monitoring and a script UTLBSTAT. SQL to terminate rnonitoring. The system catalog contains details about the sizes of tables, the distribution of values in index keys, and the like. The plan generated by the DBMS for a given query can be viewed in a graphical display that shows the estimated cost for each plan operator. While the details are specific to each vendor, all InajaI' DBMS products on the market today provide a suite of such tools.

To create a good physical database design and tune the systenl for performance in response to evolving user requirelnents, a designer HUlst understand the workings of a DBMS, especially the indexing and query processing techniques supported by the DBMS. If the database is expected to be accessed concurrently by rnany users, or is a *distributed database,* the task beeornes lnore cornplicated and other features of a, DBl\1S CaIne into play. We discuss the ilnpact of concurrency on database design in Section 20.10 and distributed databases in Chapter 22.

20.1.1 Database Workloads

The key to good physical design is arriving at an accurate description of the expectedworkloa.d. A workload description includes the follCJ\ving:

1. A list of queries (with their frequency, as a ratio of all queries / updates).
2. A list of updates and their frequencies.
3. Performance goals for each type of query and update.

For each quer.y in the workload, we HUlst identify

- Which relations are accessed.
- \Vhich attributes are retained (in the SELECT clause).

- \Vhich attributes have selection Of join conditions expressed on thern (in the WHERE clause) and how selective these conditions are likely to be.

SiInilarly, for each update in the \vorkloacl, we Blust identify

- Which attributes have selection or join conditions expressed on therll (in the WHERE clause) and how selective these conditions are likely to be.
- B The type of update (INSERT, DELETE, or UPDATE) and the updated relation.
- For UPDATE cOHnuands, the fields that are rnodified by the update.

R.elllleluber that queries and updates typically have parameters, for exarnple, a debit or credit operation involves a particular account nUlnber. rrhe values of these paralneters deterlnine selectivity of selection and join conditions.

Updates have a query cornponent that is used to find the target tuples. This cOlllponent can benefit froIn a good physical design and the presence of indexes. On the other hand, updates typically require additional work to ITlaintain indexes on the attributes that they 111odify. Thus, while queries can only benefit froill the presence of an index, an index rnay either speed up or slow down a given update. Designers should keep this trade-off in rnind when creating indexes.

20.1.2 Physical Design and Tuning Decisions

Irnportant decisions rnade during physical database design and database tuning include the followving:

1. Choice of indexes to create:
   - Which relations to index and which field or cornbination of fields to choose as index search keys.
   - For each index, should it be clustered or ul1clustered?
2. *Tuning the conceptual schema:*
   - *Alternative normalized schemas:* We usually have rnore than one way to decompose a schclua into a desired IlOl'Inal fOITn (BCNF or 3NF). A choice can be rnade on the basis of perforrnance criteria.
   - *Denormalization:* We might want to reconsider scherna decolnposibons ca.rried Olll for norrnalization. during the conceptual schern.a design process to irnprove the perforrnance of queries that involve attributes fr0111 several previously decornposed relations.

- *Vertical partitioning:* Under certain circurnstances we rnight 'want to further decornpose relations to ilnprove the perfornlance of queries that involve only a few attributes.
- *Views:* We luight 'want to add sorne views to nlask the changes in the conceptual scherna fr0111 users.

3. *Query and transaction tuning:* Frequently executed queries and transactions ulight be rewritten to run faster.

In parallel or distributed databases, which we discuss in Chapter 22, there are additional choices to consider, such as whether to partition a relation across different sites or whether to store copies of a relation at multiple sites.

20.1.3 Need for Database Thning

Accurate, detailed workload infonnation 111Cly be hard to corne by while doing the initial design of the systen1. Consequently, tuning a database after it has been designed and deployed is ilnportant---we HlllSt refine the initial design in the light of actual usage patterns to obtain the best possible perfonnance.

The distinction bet\veen database design and database tuning is soruewhat arbitrary. We could consider the design process to be over once an initial conceptual sche1na is designed and a set of indexing and clustering decisions is nlade. Any subsequent changes to the conceptual scherna or the indexes, say, would then be regarded as tuning. Alternatively, we could consider sorne refinernent of the conceptual scheula (and physical design decisions afl'ected by this refinernent) to be part of the physical design process.

Where we draw the line between design and tuning is not very irnpoItant, and we sirnply discuss the issues of index selection and database tuning without regard to when the tuning is carrier} out.

20.2 GUIDELINES FOR INDEX SELECTION

In considering which indexes to create, we begin with the list of queries (including queries tha,t a.ppear as part of update operations). ()bviously, only relations accessed by some query need to be considered as candidates for indexing, and the choice of attributes to index is guided by the conditions that appear in the WHERE clauses of the queries in the \vorkload. The presence of suitable indexes can significantly irnprove the evaluation plan for (1, query, as we saw in Chapters 8 and 12.

One approach to index selection is to consider the Ulost irnportant queries in turn, and, for each, deterrnine \vhich plan the optimizer would choose given the indexes currently on our list of (to be created) indexes. Then\ve consider whether we can arrive at a substantially better plan by adding more indexes; if so, these additional indexes are candidates for inclusion in our list of indexes. In general, range retrievals benefit froIn a B+ tree index, and exact-IHatch retrievals benefit frorn a hash index. Clustering benefits range queries, and it benefits exact-rnatch queries if several data entries contain the saIne key value.

Before adding an index to the list, however, we lnust consider the impact of having this index on the upda,tes in our workload. As we noted earlier, although an index can speed up the query cornponent of an update, all indexes on an updated attribute---{)n *any* attribute, in the case of inserts and deleteslnust be updated whenever the value of the attribute is changed. Therefore, we must sOlnetirnes consider the trade-off of slo\ving sorne update operations in the workload in order to speed up some queries.

Clearly, choosing a good set of indexes for a given workload requires an understanding of the available indexing techniques, and of the workings of the query optiruizer. The following guidelines for index selection sunnnarize our discussion:

Whether to Index (Guideline 1): The obviollS points are often the lnost important. Do not build an index unless sorne query including the query cOlnponents of updates benefits frolu it. Whenever possible, choose indexes that speed up rIlore than one query.

Choice of Search Key (Guideline 2): Attributes rnentioned m a, WHERE clause are ca,ndidates for indexing.

- An exact-match selection condition suggests that we consider an irldex on the selected attributes, ideally, a hash index.
- *j*\ range selection condition suggests that we consider a 13+- tree (Of ISAM) index on the selected attrilnltes. *j*\ B+ tree index is usually preferable to an ISAM index. A.n ISAM irlclex rnay be worth considering if the relation is infrequently updated, but we assume that a B+ tree index is always chosen over an ISAM index, for sirnplicity.

Multi-Attribute Search :Keys (Guideline 3): Indexes with multiple-attribute search keys slH)uld be considered in the follc)\ving two situ<ltion.s:

- *j*\ WHERE clause includes conditinns 011 Inore t.han one attribute of a relation.

- They enable index-only evaluation strategies (i.e., accessing the relation can be avoided) for important queries. (This situation Gould lead to attributes being in the search key even if they do not appear in WHERE clauses.)

When creating indexes on search keys with rnultiple attributes, if range queries are expected, be careful to order the attributes in the search key to match the quenes.

Whether to Cluster (Guideline 4): At lllost one index on a given relation can be clustered, and clustering affects perfonnance greatly; so the choice of clustered index is iInportant.

- As a rule of tlnunb, range queries are likely to benefit the 1110St froIll clustering. If several ra,nge queries are posed on a relation, involving different sets of attributes, consider the selectivity of the queries and their relative frequency in the workload when deciding which index should be clustered.
- If an index enables an index-only evaluation strategy for the query it is intended to speed up, the index need not be clustered. (Clustering Inatters only when the index is used to retrieve tuples fr(nll the underlying relation.)

Hash versus Tree Index (Guideline 5): A B+ tree index is usually preferable because it supports range queries as well as equality queries. A hash index is better in the following situations:

- The index is intended to support index nested loops join; the indexed relation is the inner relation, and the search key includes the join colurllns. In this case, the slight ilnproveIllent of a hash index over a B+ tree for equality selections is rnagnified, because an equality selection is generated for each tuple in the outer relation..
- rrllcre is a very important equality query, and no range queries, involving the sea.rch key attributes.

Balancing the Cost of Index Maintenance (Guideline 6): After drawing up a 'wishlist' of indexes to create, consider the irnpact of each index on the updates in the workload.

- If maintaining an index slows down frequent update operations, consider dropping the index.
- Keep ill mind, however, that adding an index Illay 'well speed up a given update operation. For exanlplc, an index on employee IDs could speed up the operation of increasing the salary of a given ernployee (specified by ID).

20.3 BASIC EXAMPLES OF INDEX SELECTION

The follawing examples illustrate how to choose indexes during database design, continuing the discussion froln Chapter 8, where we focused on index selection for single-table queries. The schernas used in the exarnples are not described in detail; in general, they contain the attributes nalned in the queries. Additional inforlnation is presented when necessary.

Let us begin with a silnple query:

```
SELECT  E.enaIne, D.rugI'
FROM    Enlployees E, Departments D
WHERE   D.dname='Toy' AND E.dno=D.dno
```

The relations rnentioned in the query are Enlployees and Departnlents, and both conditions in the WHERE clause involve equalities. Our guidelines suggest that we should build hash indexes on the attributes involved. It seeIns clear that we should build a hash index on the *dnaTne* attribute of Departments. But consider the equality *E. dno=D. dno.* Should we build an index (hash, of course) on the *dno* attribute of Departrnents or Ernployees (or both)? Intuitively, we want to retrieve Departments tuples using the index on *dnarne* because few tuples are likely to satisfy the equality selection *.D. dnaTne= 'Toy'.*[1] For each qualifying Departrnents tuple, we then find lnatching EInployees tuples by using an index on the *dno* attribute of Ernployees. So, we should build an index on the *dno* field of En1ployees. (Note that nothing is gained by building an additional index on the *dno* field of Departrnents because Departnlents tuples are retrieved using the *dna:rne* index.)

Our choice of indexes was guided by the query evaluation plan we wanted to utilize. This consideration of a, potential evaluation plan is connnon while rnaking physical design decisions. U·nderstanding query optirnization is very useful for physical design. We show the desired plan for this query in Figure 20.1.

As a variant of this query, suppose that the WHERE clause is rnodified to be WHERE *J). dnarne='Toy'* AND E'.*dno=D. dno* AND *E.age=25.* Let us consider alternative evaluation plans. ()ne good plan is to retrieve Departrnents tuples that satisfy the selection on *dnarne* and retrieve rnatching Ernployees tuples by using an index on the *dno* field; the selection on *age* is then applied on-the-fly. However, unlike the previous variant of this query, we do not really need to have an index on the *dna* field of Ernployees if we have an index. on *age.* In this

[1] This is only a heuristic. If *dname* is not the key, and we have no statistics to verify this cla.inl. it is possible that several tuples satisfy this condition.

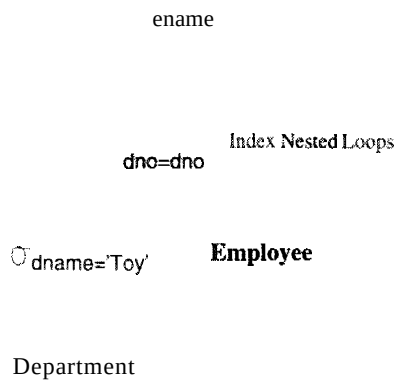


Figure 20.1 A Desirable Query Evaluation Plan

case we can retrieve Departrnents tuples that satisfy the selection on *dnarrte* (by using the index on *dname,* as before), retrieve Ernployees tuples that satisfy the selection on *age* by using the index on *age,* and join these sets of tuples. Since the sets of tuples we join are srnall, they fit in 111emory and the join Inethod is unirnportant. This plan is likely to be sornewhat poorer than using an index on *dno,* but it is a reasonable alternative. rrherefore, if we have an index on *age* already (prolnpted by sorne other query in the workload), this variant of the sarnple query does not justify creating an index on the *dno* field of Ernployees.

Our next query involves a range selection:

```
SELECT  E.enan1e, I),dnarne
FROM    Elnployees E, Departrnents D
WHERE   B.sal BETWEEN 10000 AND 20000
        AND E.hobby='Starups' AND E.dno=D.dno
```

This query illustrates the use of the BETWEEN operator for expreSSIng range selections. It is equivalent to the condition:

$$10000 \leq \text{E.sal AND E.sal} \leq 20000$$

The use of BETWEEN to express rarlge conditions is reconunended; it lnakes it easier for both the user and the optilnizer to recognize both parts of the range selection.

Returning to the exarnple query, both (nonjoin) selections are on the Ernployees relation. Therefore, it is clear that a plan in which Eluployees is the outer relation and I)epartrnents is the inner relation is the best, as in the previous query, and we should build a hash index on the *dno* attribute of Departlnents. But which index should we build on Ernployees? A B+ tree index on. the *sal* attribute would help with the range selection, especially if it is clustered. A

hash index on the *hobby* attribute would help ·with the equality selection. If one of these indexes is available, we could retrieve Ernployees tuples using this index, retrieve rnatching Departruents tuples using the index on *dno*, aJld apply all rernaining selections and projections on-the-fly. If both indexes are available, the optirnizer would choose the rno1'e selective index for the given query; that is, it \vollld consider which selection (the range condition on *salary* or the equality on *hobby)* has fe\ver qualifying tuples. In general, which index is rnore selective depends on the data. If there are very few people with salaries in the given range and rnany people collect starnps, the B-t- tree index is best. Otherwise, the hash index on *hobby* is best.

If the query constants are known (as in our exarnple), the selectivities can be estiInated if statistics on the data are available. Otherwise, as a rule of thurnb, an equality selection is likely to be rnore selective, and a reasonable decision would be to create a hash index on *hobby.* Sornethnes, the query constants are not known—we rnight obtain a query by expanding a query on a view at rUIl-tirrle, or we rnight have a query in Dynalnic SQL, which allows constants to be specified as *wild-card variables* (e.g., *%X)* and instantiated at run-tinle (see Sections 6.1.3 and 6.2). In this case, if the query is very important, we lllight choose to create a B+ tree index on *sal* and a hash index on *hobby* and leave the choice to be rnade by the optirnizer at run-tirrle.

20.4 CLUSrrERING AND INDEXING

Clustered indexes can be especially iInportant while accessing the inner relation in an index nested loops join. To understand the relationship between clustered hldexes and joins, let us revisit our first exarnple:

```
SELECT  E.enanle, D.rngr
FROM    Employees E, I)epartrnentsD
WHERE   D.dname='Toy' AND E.dno=D.dno
```

We concluded that a good evaluation plan is to use an index on *dname* to retrieve Departments tuples satisfying the condition on *dnarne* and to find. rnatching Ernployees tuples using an index on *dna.* Should these indexes be clustered? G·iven our asslunption that the number of tuples satisfying 1). *dname='Toy'* is likely to be small, we should build an unclustered index on *dname.* ()n the other hand, Employees is the inner relation in an index nested loops join and *dna* is not a candidate key. This situation is a strong argument that the index on the *dno* field of Ernployees 8ho111(1 be clustered. In fact, because the join consists of repeatedly posing equality selections on the dnofield of the inner relation, this type of quer,Y is a stronger justification for rnaking the index on *dno* clustered than a sirnple selection query such as the previous selection on

hobby. (Of courso, factors such as selectivities and frequency of queries have to be taken into account as well.)

'rhe following oxaluplc, very sirnilar to the previous one, illustrates how clustered indexes can be used for sort-rnerge joins:

```
SELECT  E.enarne,D.rngr
FROM    Ernployees E, DepartlTlents D
WHERE   E.hobby='Starnps' AND E.dno=D.dno
```

This query differs froIll the previous query in that the condition *E.hobby='Starnps'* replaces *D.dname='Toy'*. Based on the assumption that there are few ernployees in the Toy departrnent, we chose indexes that would facilitate an indexed nested loops join with DepartlTlents as the outer relation. Now, let us suppose that rllany ernployees collect stamps. In this case, a block nested loops or sort-rnerge join Blight be rnore efficient. A sort-rnerge join can take advantage of a clustered B+ tree index on the *dno* attribute in Departrnents to retrieve tuples and thereby avoid sorting Departrnents. Note that an unelustered index is not useful----since all tuples are retrieved, performing one I/O per tuple is likely to be prohibitively expensive. If there is no index on the *dno* field of Ernployees, we could retrieve Ernployees tuples (possibly using an index on *hobby,* especially if the index is clustered), apply the selection *E.hobby='Starnps'* on-the-fly, and sort the qualifying tuples on *dno.*

As our discllssion has indicated, when we retrieve tuples using an index, the irnpact of clustering depends on the rnunber of retrieved tuples, that is, the nuruber of tuples that satisfy the selection conditions that rnatch the index. An unclustered index is just as good as a clustered index for a selection that retrieves a single tuple (e.g., an equality selection on a candidate key). As the llurnber of retrieved tuples increases, the unclustered index quickly becoHlcs rnore expensive than e'ven a sequential scan of the entire relation. Although the sequential scan retrieves all tuples, each page is retrieved exactly once, whereas a page rllay be retrieved as often as the rnunber of tuples it contains if an unclustered index is usee!' If blocked l/C) is perforrned (as is COl1nnon), the relative advantage of sequential scan versus an llnclustered index increases further. (Blocked T/C) also speeds up access using a clustered index, of course.)

We illustrate the relationship between the number of retrieved tuples, viewed as a percentage of the total nurnber of tuples in the relation, and the cost of various access rnethods in .Figure 20.2. We assume that the query is a selection on a single relation, for sirnplicity. (Note that this figure reflects the C()st of writing out the result; otherwise, the line for seqnential scan weHlld be flat.)

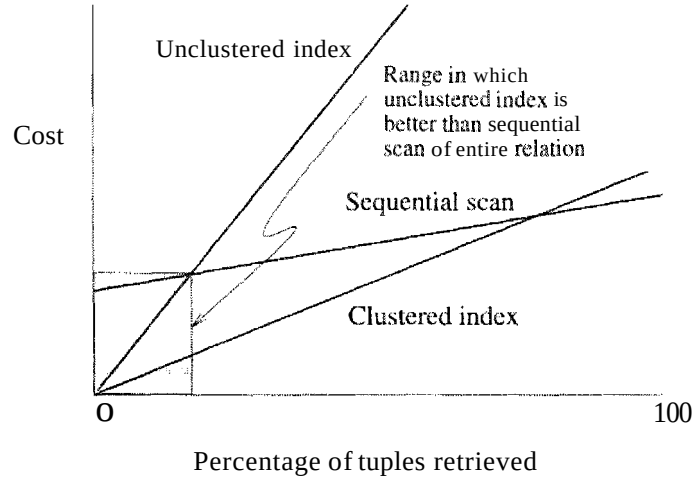


Figure 20.2 The Impact of Clustering

20.4.1 Co-clustering Two Relations

In our description of a typical database system architecture in Chapter g, we explained how a relation is stored as a file of records. Although a file usually contains only the records of SOIn8 one relation, SCHue systeIlls allow records frorn Inore than one relation to be stored in a single file. rrhe database user can request that the records froIll two relations be interleaved physically in this 111anne1'. This data layout is sornetiInes referred to as co-clustering the two relations. We now discuss when co-clustering can be beneficial.

As an exarnple, consider two relations with the following schernas:

Parts(pid: integer, *pnarne:* string, *cost:* integer, *8upplierid:* integer)
Assembly(*partid:* integer, *componentid:* integer, *quantity:* integer)

In this scherna the *componentid* field of Assernbly is intended to be the *pid* of sorne part that is used as a cornponent in assernbling the part with *pid* equal to *partid.* Therefore, the Assernbly table represents a 1:N relationship between parts and their subparts; a part can have rnany s11bparts, but each part is the subpart of at rnost one part. In, the Parts table, *pid* is the key. For cOlnposite parts (those assernbled frorIl other parts, as indicated by the contents of Assclnb1y), the *cost* field is taken to be the cost of assembling the part frorn its subparts.

Suppose that a frequent query is to find the (inllnediate) subparts of all parts supplied by a given supplier:

```
SELECT  P.piel, .A.componentid
FROM    Parts J') Assembly A
```

```
WHERE   P.piel = A.partid AND P.supplierid = 'Acme'
```

A good evaluation plan is to apply the selection condition on Parts and then retrieve rnatching Asselnbly tuples throngh an index on the *partid* field. Ideally, the index on *partid* should be clustered. rrhis plan is reasonably good. :However, if such selections are COHU110n and we want to optirnize thorn further, we can *co-cluster* the two tables. In this approach, we store records of the two tables together, \vith each Parts record *P* follc)\ved by all the Assernbly records *A* such that *P.pid = A.partid*. This approach improves on storing the two relations separately and having a, clustered index on *paTtid* because it does not need an index lookup to find the A.ssernbly records that rnatch a given Parts record. Thus, for each selection query, we save a few (typically two or three) index page l/Os.

If we are interested in finding the imrnediate subparts of *all* parts (i.e., the preceding query with no selection on *supplierid)*, creating a clustered index on *partid* and doing an index nested loops join with Assembly as the inner relation offers good perfonnance. An even better strategy is to create a clustered index on the *paTtid* field of Assernbly and the *pid* field of Parts, then do a sort-rnerge join, using the indexes to retrieve tuples in sorted order. This strategy is comparable to doing the join using a co-clustered organization, which involves just one scan of the set of tuples (of Parts and Asselnbly, which are stored together in interleaved fashion).

The real benefit of co-clustering is illustrated by the following query:

```
SELECT  P.pid,A.componentid
FROM    Parts P, Assernbly A
WHERE   P.pid = A.partid AND P.cost=10
```

Suppose that rnany parts have *cost* = 10. This query essentially a,rnonnts to a collection of queries in which we are given a Parts record and want to find rnatching Assernbly records. If we have an index on the *cost* field of Parts, we can retrieve qualifying Parts tuples. For each such tuple, we have to use the index on Assernbly to locate records with the given *pid*. rrhe index access for A.ssernbly is avoided if we have a co-clustered organization. (()f courS8, we still require aII index on the *cost* attribute of Parts tuples.)

Such an optirnization is especially irnportant if we want to traverse several levels of the part-subpart hierarchy. .For example, a COnll110Il query is to find the totaJ cost of a part, which requires us to repeatedly carry out joins of Pa,rts and Asscrnbly. Incidentally, if we do not know the nurnber of levels in the hierarchy iII advance, the nUlnber of joins varies and the query cannot be expressed in SQL. The query can be answered by ernbedcling an SQL statement

for the join insicie an iterative host language prograrll. How to express the query is orthogonal to our lnain point here, which is that co-clustering is especially beneficial when the join in question is carried out very frequently (either because it arises repeatedly in an important query such as finding total cost, or because the join query itself is asked frequently).

To sUIlunarize co-clustering:

- It can speed up joins, in particular key·foreign key joins corresponding to l:N relationships.
- A sequential scan of either relation becornes slower. (In our exalnple, since several Assenlbly tuples are stored in between consecutive Parts tuples, a scan of all Parts tuples becornes slower than if Parts tuples were stored separately. SiInilarly, a sequential scan of all Assernbly tuples is also slower.)
- All inserts, deletes, and updates that alter record lengths becorne slower, thanks to the overheads involved in ruaintaining the clustering. (We do not discuss the irnplernentation issues involved in co-clustering.)

20.5 INDEXES THAT ENABLE INDEX-ONLY PLANS

This section considers a nU111ber of queries for which we can find efficient plans that avoid retrieving tuples froln one of the referenced relations; instead, these plans scan an associated index (which is likely to be lnuch srnaller). An index that is used (only) for index-only scans does *not* have to be clustered because tuples fronl the indexed relation are not retrieved.

This query retrieves the lnanagers of depal'truents with at least one ernployee:

```
SELECTD.rugr
FROM    Departrnents I), Employees E
WHERE   I).dno=E.dno
```

()bserve that no attributes of Ernployees are retained. If we have an index on the *dno* field of Employees, tl1e optirnization of doirlg an index nested loops join using an index-onl:y searl for the inner relation is applicable. Given this variant of the query, the correct decision is to build an uncIustered index on the *dna* field of Employees, rather thaIl a clustered index.

rrhe next query takes this idea a step further:

```
SELECT  ]).rngr, E.eid
FROM    ])epartrnents I), Employees E
WHERE   D.dno=E.dno
```

If we have an index on the *dno* field of Employees, we can use it to retrieve Employees tuples during the join (\vith Departrnents as the outer relation), but unless the index is clustered, this approach is not be efficient. ()n the other hand, suppose that we have a B+- tree index on *(dna, e'id).Now* all the inforrnation we need about an Employees tuple is contained in the data entry for this tuple in the index. We can use the index to find the first data entry \vith a given *elno;* all data entries 'with the same *dno* are stored together in the index. (Note that a hash index on the cOlnposite key *(dna, eid)* cannot be used to locate an entry with just a given *dno!)* We can therefore evaluate this query using an index nested loops join with Departlnents as the outer relation and an index-only scan of the inner relation.

20.6 TOOLS TO ASSIST IN INDEX SELEC"fION

The rUllnber of possible indexes to consider building is potentially very large: For each relation, we can potentially consider all possible subsets of attributes as an index key; we have to decide on the ordering of the attributes in the index; and we also have to decide which indexes should be clustered and which unclustered. Many large applications---for exalnple enterprise resource planning systems—create tens of thousands of different relations, and rnanual tuning of such a large schelna is a daunting endeavor.

The difficulty and irnportance of the index selection task rnotivated the development of tools that help database adrninistrators select appropriate indexes for a given workload. The first generation of such index tuning wizards, or index advisors, were separate tools outside the database engine; they suggested indexes to build, given a workload of SQL queries. rfhe rnain drawback of these systerns was that they had to replicate the database query optirnizer's cost rnodel in the tuning tool to rnake sure that the optimizer would choose the sanlC query evaluation plans as the design tool. Since query optirnizers cha,nge froIn release to release of a conunercial database systern, considerable effort was needed to keep the tuning tool and the database optirnizer synchronized. The rnost recent generation of tuning tools are integrated \vith the database engine and use the database query optiluizer to estimate the cost of a workload given a set of indexes, avoiding duplication of the query optirnizer's cost rnodel into an external tool.

20.6.1 Automatic Index Selection

\{Ve call a set of indexes for a given database scherna. an index configuration. We assume that a, query workload is a set of queries over a database scherna where each query has a frequency of occurrence assigned to it. Given a database schelna and a, workload, the cost of an index configuration is the expected

cost of running the queries in the workload given the index configuration taking the different frequencies of queries in the workload into account. Given a database schema and a query workload, we can now define the problel11 of automatic index selection as finding an index configuration with nlinirnal cost. A.s in query optinlization, in practice our goaJ is to find a *good* index configuration rather than the true optirnal configuration.

Why is automatic index selection a hard problern? Let us calculate the nUlnber of different indexes with c attributes, assurning that the table has n attributes. For the first attribute in the index, there are n choices, for the second attribute $n-1$, and thus for a c attribute index, there are overall $n \cdot (n-1) \ldots (n-c+1) = \frac{n!}{(n-c)!}$ different indexes possible. The total nurnber of different indexes with up to c attributes is

$$\sum_{i=1}^{c} \frac{n!}{(n-i)!}$$

For a table with 10 attributes, there are 10 different one-attribute indexes, 90 different two-attribute indexes, and 30240 different five-attribute indexes. For a cornplex workload involving hundreds of tables, the nurnber of possible index configurations is clearly very large.

The efficiency of autornatic index selection tools can be separated into two components: (1) the nurnber of candidate index configurations considered, and (2) the nurnber of optimizer calls necessary to evaluate the cost for a configuration. Note that reducing the search space of candidate indexes is analogous to restricting the search space of the query optiInizer to left-deep plans. In lnany cases, the optirnal plan is not left-deep, but alllong all left-deep plans there is usually a plan whose cost is close to the optirnal plan.

We can easily reduce the time taken for autornatic index selection by reducing the nUlnber of candidate index configurations, but the srnaller the space of index. configurations considered, the farther away the final index configuration is [1'0111 the optirnal index configllration. rrherefore, different index tuning \vizards prune the search space differently, for exarnple, by considering only one- or two-attribute indexes.

20.6.2 How Do Index Thning Wizards Work?

All index tuning \vizards search a set of candidate indexes for an index configuration with lowest cost. Tools differ in the space of candidate index configurations they consider and how they search this space. We describe one representative algoritlun; existing tools implement 'variants of this algorithrn, hut their irnplernentations have the sanIe basic structure.

The DB2 Index Advisor. The DB2 Index Advisor is a tool for automatic index recommendation given a workload. The workload is stored in the database system in a table called ADVISE_WORKLOAD. It is populated either (1) by SQL statements in the DB2 dynamic SQL statement cache, a cache for recently executed SQL statements, (2) with SQL statements from packages---groups of statically compiled SQL statements, or (3) with SQL statements from an online monitor called the Query Patroller. The DB2 Advisor allows the user to specify the maximum amount of disk space for new indexes and a maximum time for the computation of the recommended index configuration.

The DB2 Index Advisor consists of a program that intelligently searches a subset of index configurations. Given a candidate configuration, it calls the query optimizer for each query in the ADVISE_WORKLOAD table first in the RECOMMEND_INDEXES mode, where the optimizer recommends a set of indexes and stores them in the ADVISE_INDEXES table. In the EVALUATE_INDEXES mode, the optimizer evaluates the benefit of the index configuration for each query in the ADVISE-WORKLOAD table. The output of the index tuning step is are SQL DDL statements whose execution creates the recommended indexes.

The Microsoft SQL Server 2000 Index Tuning Wizard. Microsoft pioneered the implementation of a tuning wizard integrated with the database query optimizer. The Microsoft Tuning Wizard has three tuning modes that permit the user to trade off running time of the analysis and number of candidate index configurations examined: *fast*, *medium*, and *thorough*, with *fast* having the lowest running time and *thorough* examining the largest number of configurations. To further reduce the running time, the tool has a sampling mode in which the tuning wizard randomly samples queries from the input workload to speed up analysis. Other parameters include the maximum space allowed for the recommended indexes, the maximum number of attributes per index considered, and the tables on which indexes can be generated. The Microsoft Index Tuning Wizard also permits *table scaling*, where the user can specify an anticipated number of records for the tables involved in the workload. This allows users to plan for future growth of the tables.

Before we describe the index tuning algoriUlIn, let us consider the problell1 of estiInating the cost of a configuration. Note that it is not feasible to actually create the set of indexes in a candidate configuration and then optirnize the query workload given the physical index configuration. Creation of even a single candidate configuration with several indexes lIlight take hours for large databases and put considerable load on the database systerIl itself. Since we want to exauline a large nUlnber of possible candidate configurations, this approach is not feasible.

Therefore index tuning algorithrIls usually *simulate* the effect of indexes in a candidate configuration (unless such indexes already exist). Such what-if indexes look to the query optilIlizer like any other index and are taken into account when calculating the cost of the workload for a given configuration, but the creation of what-if indexes does not incur the overhead of actual index creation. Commercial database systelIls that support index tuning wizards using the database query optirnizer have been extended with a module that permits the creation and deletion of what-if indexes with the necessary statistics about the indexes (that are used when estirnating the cost of a query plan).

We now describe a representative index tuning algorithm. The algorithm proceeds in two steps, *candidate index selection* and *configuration enumeration.* In the first step, we select a set of candidate indexes to consider during the second step as building blocks for index configurations. Let us discuss these two steps in Inore detail.

Candidate Index Selection

We saw in the previous section that it is iInpossible to consider every possible index, due to the huge nUluber of candidate indexes available for larger database schernas. ()ne heuristic to prune the large space of possible indexes is to first tune each query in the workload independently and then select the union of the indexes selected in this first step as input to the second step.

:For a query, let us introduce the notion of an indexable attribute, which is an attribute whose appearance in an index could change the cost of the query. An indexable attribute is an attribute on \vhich the WHERE-part of the query has a condition (e.g., an equality predicate) or the attribute appears in a GROUP BY or ORDER BY clause of the SQL query. An admissible index for a query is an index that contains only indexable attributes in the query.

How do we select candidate indexes for an individual query? ()ne approach is a basic enumeration of all indexes with up to k attributes. \Ve start \ivith all indexable attributes as single attribute candidate indexes, then add all com-

binations of two indexable attributes as candidate indexes, and repeat this procedure until a user-defined size threshold *k*. This procedure is obviously very expensive as we add overall $n + n \cdot (n - 1) + \ldots + n \cdot (n - 1) \ldots (n - k + 1)$ candidate indexes, but it guarantees that the best index with up to *k* attributes is all10ng the candidate indexes. The references at the end of this chapter contain pointers to faster (but less exhaustive) heuristieal search algorithrns.

Enumerating Index Configurations

In the second phase, we use the candidate indexes to enUInerate index configurations. As in the first phase, we can exhaustively enurnerate all index configurations up to size *k,* this time cornbining candidate indexes. As in the previous phase, more sophisticated search strategies are possible that cut down the number of configurations considered while still generating a final configuration of high quality (i.e., low execution cost for the final workload).

20.7 OVERVIEW OF DATABASE TUNING

After the initial phase of database design, actual use of the database provides a valuable source of detailed information that can be used to refine the initial design. Many of the original assumptions about the expected workload can be replaced by observed usage patterns; in general, some of the initial workload specification is validated, and some of it turns out to be wrong. Initial guesses about the size of data can be replaced with actual statistics frorn the systern catalogs (although this inforrnation keeps changing as the systern evolves). Carefulrnonitoring of queries can reveal unexpected problerlls; for example, the optimizer lllay not be using SOIne indexes as intended to produce good plans.

Continued database tuning is irnportant to get the best possible perforrnance. In this section, we introduce three kinds of tuning: *tuning indexes, tuning the conceptual scherna,* and *tuning queries.* OUr discussion of index selection also applies to index tuning decisions. Conceptual schcrna and query tuning are discussed further in Sections 20.8 and 20.9.

20.7.1 Thning Indexes

The initial choice of indexes rnay be refined for one of several reasons. The sirnplest reason is that the observed \vorkload reveals that scnne queries and updates considered irnportant in the initial\vorkload specification are not very frequent. The observed workload rnay a1so identify SCHne new queries and updates that *are* inlportant.The initial choice of indexes has to be reviewed in light of this new inforrnation. Scnne of the original in,dexes rnay be dropped and

new ones added. The reasoning involved is siInilar to that used in the initial design.

It Inay also be discovered that the optimizer in a given systenl is not finding some of the plans that it was expected to. For exaluple, consider the following query, which we discussed earlier:

```
SELECT  D.Ingr
FROM    Ernployees E, Departulents D
WHERE   D.dnamc='Toy' AND E.dno=D.dno
```

A good plan here would be to use an index on *dnarne* to retrieve Departnlents tuples with *dnarne= 'Toy'* and to use an index on the *dno* field of Employees as the inner relation, using an index-only scan. Anticipating that the optirnizer would find such a plan, we rnight have created an unclustered index on the *dno* field of Ernployees.

Now suppose queries of this fonn take an unexpectedly long time to execute. We can ask to see the plan produced by the optiInizer. (Most commercial systerIls provide a simple cOillrnand to do this.) If the plan indicates that an index-only scan is not being used, but that Employees tuples are being retrieved, we have to rethink our initial choice of index, given this revelation about our system's (unfortunate) lhnitations. An alternative to consider here would be to drop the unclustered index on the *dno* field of EUlployees and replace it with a clustered index.

SOUle other COllllnon lirnitations of optiInizers are that they do not handle selections involving string expressions, arithrnetic, or *null* values effectively. We discuss these points further when we consider query tuning in Section 20.9.

In addition to re-exarnining our choice of indexes, it pays to periodically reorganize S(Hne indexes. For example, a static index, such as an ISAM index, Illay have developed long overflow chains. Dropping the index and rebuilding it—if feasible, given the interrupted access to the indexed relation—can substantially irnprove access tiTHes through this index. Even for a dynarnic structure such as a 13+ tree, if the implernentation does not rnerge pages on deletes, space occupancy can decrease considerably in SaIne situations. This in turn rnakes the size of the index (in pages) larger than necessary, and could increase the height and therefore the access tilne. Ilebuilding the index should be considered.Extensive updates to a clustered index rnight also lead to overflow pages being allocated, thereby decreasing the degree of clustering. Again, rehuilding the index Inay be vvorthwhile.

Finally, note that the query optinlizer relies on statistics rnaintained in the SystCIll catalogs. These statistics are updated only when a special utility progranl is run; be sure to run the utility frequently enough to keep the statistics reasonably current.

20.7.2 Thning the Conceptual Schema

In the course of database design, we rnay realize that our current choice of relation SChelllaS does not enable us rneet our perforrnance objectives for the given workload with any (feasible) set of physical design choices. If so, we 11lay have to redesign our conceptual scherna (and re-exarnine physical design decisions affected by the changes we rnake).

We rnay realize that a redesign is necessary during the initial design process or later, after the systern has been in use for a while. Once a database has been designed and populated with tuples, changing the conceptual scherna requires a significant effort in tenrlS of rnapping the contents of the relations affected. Nonetheless, it rnay be necessary to revise the conceptual scherna in light of experience with the systern. (Such changes to the schema of an operational systerll are sometirnes referred to as schema evolution.) We now consider the issues involved in conceptual scherna (re)design frorn the point of vie\v of perforrnance.

rrhe rnain point to understand is that *our choice of conceptual 8cherna should be guided by a cons'ideration of the queries and 'updates in our 'workload"* in addition to the issues of redundancy that rllotivate nonnalization (which we discussed in Chapter 19). Several options rnust be considered while tuning the conceptual scherna:

- We lIlay decide to settle for a 3NF design instead of a BCN'F design.
- If there are two ways to decornpose a given schelna into 3NF or BCNF, our choice should be guided by the workload.
- Sometimes we rnight decide to further decornpose a relation that is *already* rn BCNF.
- In other situations, we rnight *denormalize.* rrhat is, we rnight choose to replace a collection of relations obtained by a dec(nnposition frorn a larger relation with the original (larger) relation, even though it suffers frorn 80rne redundancy problerl1s. Alternatively, we rnight choose to add sorne fields to certain relations to speed up SCHne irnportant queries, even if this leads to a redundant storage of 80rne infonnation (anei, consequently, a scherna that is in neither 3NF nor BCNF).

- This discussion of nonnalization has concentrated on the technique of *decomposition*, 'which arnounts to vertical partitioning of a relation. Another technique to consider is *horizontal partitioning* of a relation, which \vould lead to having two relations with identical schernas. Note that we are not talking about physically partitioning the tuples of a single relation; rather, we want to create two distinct relations (possibly with different constraints and indexes on each).

Incidentally, when \ve redesign the conceptual scherna, especially if we are tuning an existing database sche1na, it is \vorth considering whether we should create views to rnask these changes fronl users for whOIII the original schcrlla is]nore natural. We discuss the choices involved in tuning the conceptual scherna in Section 20.8.

20.7.3 Thning Queries and Views

If we notice that a query is running rnuch slower than we expected, we have to exarnine the query carefully to find the problern. SaIne rewriting of the query, perhaps in conjunction with SCHne index tuning, can often fix the problern. Sirnilar tuning rnay be called for if queries on SaIne view run slower than expected. We do not discuss view tuning separately; just think of queries on views as queries in their own right (after all, queries on views are expanded to account for the view definition before being optirnized) and consider how to tune thern.

When tuning a query, the first thing to verify is that the systern uses the plan you expect it to use. Perhaps the systelll is not finding the best plan for a variety of reasons. Sorne COIllrllon situations not handled efficiently by rnany optinlizers follow:

- A selection condition involving *null* values.
- Selection conditions involving aritlunetic or string expressions or conclitions using the OR connective. For exarnple, if we have a conclitionE.age = 2*]).age in the WHERE clause, the optirnizer rnay correctly utilize an available index on $E.age$ but fail to utilize an available index on 1).age. R,eplacing the condition by $E.age/2$ = 1).age \vould reverse the situation.
- Inability to recognize a sophisticated plan such as an index-only scan for an aggregation query involving a GROUP BY clause. ()f course, virtually no optirnizer looks for plans outside the plan space described in Chapters 12 and 15, such as nonleft-deep join trees. So a good understanding of ·what an optirnizer typically does is irnportant. In addition, the rnore a:ware you are of a given systeur's strengths arrd lirnitations, the better off you arc.

If the optirnizer is not smart enough to find the best plan (using access Inethods and evaluation strategies supported by the DB.wIS), SOHIe systern.s allow users to guide the choice of a plan by providing hints to the opthnizer; for example, users rnight be able to force the use of a particular index or choose the join order and join rnethod. A user who wishes to guide optirnization in this Inanner should have a thorough understanding of both optirnizatioll and the capabilities of the given DBMS. We discuss query tuning further in Section 20.9.

20.8 CHOICES IN TUNING **THE CONCEPTUAL SCHEMA**

We now illustrate the choices involved in tuning the conceptual scheIua through several exarnples using the following schelnas:

Contracts(*cid:* integer, *s'Upplierid:* integer, *projectid:* integer,
deptid: integer, *partid:* integer, *qty:* integer, *value:* real)
Departments(*did:* integer, *budget:* real, *annualreport:* varchar)
Parts(pid: integer, *cost:* integer)
Projects(jid: integer, *rngr:* char(20))
Suppliers(*sid:* integer, *address:* char(50))

For brevity, we often use the cornrnon convention of denoting attributes by a single character and denoting relation schernas by a sequence of characters. Consider the scherna for the relation Contracts, whic.h we denote as CSJDPQV, with each letter denoting an attribute. The Ineaning of a tuple in this relation is that the contract with *cid* C is an agreernent that supplier S (with *sid* equal to *supplierid)* will supply Q iterns of part P (with *pid* equal to *partid)* to project J (with *jid* equal to *projectid*) associated with departrnent D (with *deptid* equal to *did),* and that the value V of this contract is equal to *value.*[2]

There are two known integrity constraints with respect to Contracts. A project purchases a given part using a single contract; thus, there cannnot be two distinct contracts in which the saIne project buys the saIne part. This constraint is represented using th.e FI) $JD \rightarrow C$. Also, a departrnent purchases at rnost one part frolll any given supplier. This constraint is represented llsing the FD $SD \rightarrow P$. In addition, of course, the contract ID C is a key. The rneaning of the other relations should be obvious, and we do not describe thcrn further because we focus on the Contra.cts rela.tion.

2If this schema seems cornplicated, note that real-life situations often call for considerably more cornplex schemas!

20.8.1 Settling for a Weaker Normal Form

Consider the Contracts relation. Should we deconlpose it into sHlaller relations? Let us see what norrnal fann it is in. The candidate keys for this relation are C and JP. (C is given to be a key, and tIP functionally deterrnines C.) The only nonkey dependency is 3D $\rightarrow P$, and P is a *prime* attribute because it is part of candidate key JP. rrhus, the relation is not in BCNF—because there is a nonkey dependency-.::-but it is in 3NF.

By using the dependency $SD \rightarrow P$ to guide the decornposition, we get the two schemas SDP and CSJDQV. This decornposition is lossless, but it is not dependency-preserving. However, by adding the relation schelne CJP, we obtain a lossiess-join, dependency-preserving decoruposition into BCNF. Using the guideline that such a decorllposition into BCNF is good, we might decide to replace Contracts by three relations with schernas CJP, SDP, and CSJDQV.

However, suppose that the following query is very frequently asked: Find the nurnber of copies Q of part P ordered in contract C. 'rhis query requires a join of the decornposed relations CJP and CSJDQV (or SDP and CSJDQV), whereas it can be answered directly using the relation Contracts. The added cost for this query could persuade us to settle for a 3NF design and not decompose Contracts further.

20.8.2 Denormalization

The reasons rTIotivating us to settle for a weaker norrnal forIn may lead us to take an even rnore extrerne step: deliberately introduce SOlllC redundancy. As an exarnple, consider the Contracts relation, 'which is in 3NF. Now, suppose that a frequent query is to check that the value of a contract is less than the budget of the contracting departruent. We lllight decide to add a budget field B to Contracts. Since *did* is a key for Departrnents, we now have the dependency $D \rightarrow B$ in Contracts, \vhich InCt1IIS Contracts is not in 3NF any r11orc. Nonetheless, we rnight choose to stay with this design if the rnotivating query is sufficiently irnportant. Such a decision is clearly subjective and CaInes at the cost of significant redundancy.

20.8.3 Choice of Decomposition

Consider the Contracts relation again. Several choices are possible for dealing with the redundancy in this relation:

- We can leave Contracts as it is and accept the redundancy associated \\"ith its being in :3N:F rather than .BCNF.

- We Inight decide that we want to avoid the anolIlalies resulting froIn this redundancy by deeornposing Contracts into BCNF using one of the following ruethods:
  - We have a lossless-join decornposition into PartInfo with attributes SDP and Contractlnfo \vith attributes CSJDQV. As noted previously, this decornposition is not dependency-preserving, and to rnake it so would require us to add a third relation CJP, \vhose sale purpose is to allow us to cheek the dependency $JP \rightarrow C$.
  - We could choose to replace Contracts by just Partlnfo and Contract-Info even though this decornposition is not dependency-preserving.

R,eplacing Contracts by just Partlnfo and Contractlnfo does not prevent us frorll enforcing the constraint $JP \rightarrow C$; it only makes this n10re expensive. We could create an assertion in SQL-92 to check this constraint:

```
CREATE ASSERTION checkDep
CHECK     ( NOT EXISTS
            (SELECT *
            FROM     Partlnfo PI, Contractlnfo Cl
            WHERE    PI.supplierid=CI.supplierid
                     AND PI.deptid==CI.deptid
            GROUP BY C1.projectid, PI.partid
            HAVING   COUNT (cid) > 1 ) )
```

This assertion is expensive to evaluate because it involves a join followed by a sort (to do the grouping). In cornparison, the systerll can check that JP is a prirnary key for table CJP by rnaintaining an index on JP. This difference in integrity-checking cost is the rl1otivation for dependency-preservation. On the other hand, if updates are infrequent, this increased cost Illay be acceptable; therefore, we rnight choose not to rnaintain the table C.JP (and quite likely, an index all it).

As another exarnple illustrating decornposition choices, consider the Contracts relation again, and suppose that we also have the integrity constraint that a departrnent uses a given supplier for at rnost one of its projects: $SPQ \rightarrow V$. Proceeding as before, we have a lossless-join decornposition of Contracts into SDP and CSJDQV. Alternativcly, we could begin by using the dependency $SPQ \rightarrow V$ to guide our decomposition, and replace Contracts with SPQV and CSJDPQ. We can then dec(Hnpose CSJDPQ, guided by $51D \rightarrow P$, to obtain SDP and CSJDQ.

We now have two alternative lossless-join decornpositions of Contracts into BCNF, neither of which is dependency-preserving. The first alternative is to

replace Contracts with the relations SDP and CSJDQV. The second alternative is to replace it \vith SPQV, SDP, and CSJDQ. The addition of CJP makes the second deC0111positioll (but not the first) dependency-preserving. Again, the cost of lnaintaining the three relations CJP, SPQV, and CSJDQ (versus just CSJDQV) Illay lead us to choose the first alternative. In this case, enforcing the given FDs becornes Inore expensive. We Illight consider Ilot enforcing thern, but we then risk a violation of the integrity of our data.

20.8.4 Vertical Partitioning of BCNF Relations

Suppose that we have decided to decornpose Contracts into SDP and CSJDQV. These scheruas are in BCNF, and there is no reason to decornpose thern further from a nonl1alization standpoint. However, suppose that the following queries are very frequent:

- Find the contracts held by supplier S.
- Find the contracts placed by departrnent D.

These queries rnight lead us to decompose CSJDQV into CS, CD, and CJQV. The decornposition is lossless, of course, and the two il11portant queries can be answered by exarnining 111uch slualler relations. Another reason to consider such a dec()l11position is concurrency control *hot spots*. If these queries are COIllIllon, and the rnost COIlunon updates involve changing the quantity of products (and the value) involved in contracts, the decoulposition inlproves perforrnance by reducing lock contention. Exclusive locks are now set rnostly on the CJQV table, and reads on CS and CD do not conflict with these locks.

Whenever we decornpose a relation, we have to consider which queries the decolnposition rnight adversely affect, especially if the only rnotivation for the decoInposition is iUlproved perforrnance. For exaruplc, if another illlportant query is to find the total value of contracts held by a supplier, it would involve a join of the decornposed relations CS and CJQV. In this situation, we rnight decide against the decolnposition.

20.8.5 Horizontal Decomposition

Thus far, we have essentially considered how to replace a relation with a collection of vertical decorupositions. Sornetilnes, it is worth considering whether to replace a relation with two relations that have the sa.Dle attributes as the original relation, each containin.g a subset of the tuples in tb.e original. Intuitively, this technique is useful \vhen different subsets of tuples are queried in very distinct ways.

For example, different rules lnay govern large contracts, \vhich are defined as contracts with values greater than 10,000. (Perhaps, such contracts have to be awarded through a bidding process.) This constraint could lead to a nUlnber of queries in which Contracts tuples are selected using a condition of the forIII $value > 10,000$. On8 way to approach this situation is to build a clustered B+ tree index oII the *value* field of Contracts. Alternatively, we could replace Contracts with two relations called LargeContracts and SrnallContracts, with the obvious 11leaning. If this query is the only lllotivation for the index, horizontal decornposition offers all the benefits of the index without the overhead of index maintenance. This alternative is especially attractive if other irnportant queries on Contracts also require clustered indexes (on fields other than *val'ue).*

If we replace Contracts by two relations LargeContracts and SrnallContracts, we could r1l8sk this change by defining a view called Contracts:

```
CREATE VIEW Contracts(cid, supplierid, projectid, deptid, partid, qty, value)
        AS ((SELECT *
            FROM    LargeContracts)
            UNION
            (SELECT *
            FROM    SmallContracts))
```

However, any query that deals solely with LargeContracts should be expressed directly on LargeContracts and not on the view. Expressing the query on the view Contracts with the selection condition $value > 10,000$ is equivalent to expressing the query on LargeContracts but less efficient. This point is quite general: Although we can mask changes to the conceptual scherlla by adding view definitions, users concerned about perforrnance have to be aware of the change.

As another exanlple, if Contracts had an additional field *year* and queries typically dealt with the contracts in sorne one year, we rnight choose to pa,rtition Contracts by year. ()f course, queries that involved contracts fron1 rnore than one year rnight require us to pose queries against each of the decolllposed relations.

20.9 CHOICES IN TUNING QUERIES AND VIEWS

T'he first step in tuning a query is to understand the plan used by the DBMS to evaluate the query. 8ysten1s usually provide sorne facility for identifying the plan used to evaluate a query. Once we understand the plan selected by the systelIl, we can consider how to irnprove performance. We can consider a different choice of ilHlexes or perhaps co-clustering two relations for join queries,

guided by our understanding of the old plan and a better plan that we want theDBlVIS to use. The details are sinlilar to the initial design process.

One point'worth rnaking is that before creating new indexes we should consider \vhether rewriting the query achieves acceptable results with existing indexes. For example, consider the foll()\ving query with an OR connective:

```
SELECT  E.dno
FROM    Ernployees E
WHERE   E.hobby='Stalnps' OR E.age==10
```

If \ve have indexes on both *hobby* and *age,* we can use these indexes to retrieve the necessary tuples, but an optiInizer ruight fail to recognize this opportunity. The optinlizer rnight view the conditions in the WHERE clause as a whole as not rnatching either index, do a sequential scan of Ernployees, and apply the selections on-the-fly. Suppose we rewrite the query as the union of two queries, one with the clause *WHEREE.hobby= 'Starnps"* and the other with the clause WHERE *E.agc==10.* Now each query is answered efficiently with the aid of the indexes on *hobby* and *age.*

We should also consider rewriting the query to avoid sorne expensive operations. For exalnple, including DISTINCT in the SELECT clause leads to duplicate elirnination, which can be costly. rrhus, we should ornit DISTINCT whenever possible. For exalnple, for a query on a single relation, we can ornit DISTINCT whenever either of the following conditions holds:

- We do not care about the presence of duplicates.
- rrhe attributes lllentioned in the SELECT clause include a candidate key for the relation.

SOlnetirnes a query with GROUP BY and HAVING can be replaced by a query without these clauses, thereby eliminating a sort operation. For example, consider:

```
SELECT    MIN (E.age)
FROM      Employees E
GROUP BY  E.dno
HAVING    E.dno=102
```

This quer:y is equivalent to

```
SELECT    MIN (E.age)
FROM      Erl1ployees E
WHERE     E.dno=102
```

Cornplex queries are often written in steps, using a temporary relation. We can usually rewrite such queries without the tClnporary relation to rnake thcrn run faster. Consider the following query for cornputi.ng the average salary of departrnents rnanaged by Robinson:

```
SELECT   *
INTO     Ternp
FROM     ErnployeesE, Depa:rtruents D
WHERE    E.dno=D.dno AND D.rngrnanle='Robinson'

SELECT   T.dno, AVG (T.sal)
FROM     T'clnp T
GROUP BY T.dno
```

This query can be rewritten as

```
SELECT   E.dno, AVG (E.sal)
FROM     Elnployees E, Departments D
WHERE    E.dno==D.dno AND D.rngrnarne=='llobinson'
GROUP BY E.dno
```

The rewritten query does not 111aterialize the interrnediate relation ,Ternp and is therefore likely to be faster. In fact, the optimizer may even find a very efficient index-only plan that never retrieves Ernployees tuples if there is a cornposite B+· tree index on *(d'no, sal).* This exanlple illustrates a general observation: *By rewriting queries to avoid unnecessary temporaries, we not only avoid creating the ternporary relations, we also open up rnore optimization possibilities for the optim,izer to el;plore.*

In SCHne situations, ho\vever, if the optirnizer is unable to find a good plan for a cornplex query (typically a nested query with correlation), it rnay be worthwhile to rewrite the query using tenlporary relations to guide the optirnizer toward a good plan.

In fact, nested queries are a conunon source of inefficiency because luany optirnizers deal poorly with theIn, as discussed in Section 15.5.v'Vllenever possible, it is better to rewrite a nested query \vithout nesting and a correlated query without correlation. As already noted, a good reforrIlulation of the query rnay require us to introduce new, ternporary relations, and techniques to do so systenlatically (ideally, to be done by the optirnizer) have been \videly studied. ()ften tllough, it is possible to re\vrite nested queries ,vithout nesting or the use of ternpora,ry relations, as illustrated in Section 15.5.

20.10 IMPACT OF CONCURRENCY

In a system with IIlallY concurrent users, several additional points IllUSt be considered. Transactions obtain *locks* on the pages they access, and other transactions Ina)' be blocked waiting for locks on objects they wish to access.

We observed in Section 16.5 that blocking delays 111ust be IniniInized for good perforrnance and identified two specific ways to reduce blocking:

- Reducing the time that transactions hold locks.
- R,edllcing hot spots.

We now discuss techniques for achieving these goals.

20.10.1 Reducing Lock Durations

Delay Lock Requests: Tune transactions by writing to local prograrn variables and deferring changes to the database until the end of the transaction. This delays the acquisition of the corresponding locks and reduces the time the locks are held.

Make Transactions Faster: The sooner a transaction c01npletes, the sooner its locks are released. We have already discussed several ways to speed up queries and updates (e.g., tUllillg indexes, rewriting queries). In addition, a careful partitioning of the tuples in a relation and its associated indexes across a collection of disks can significantly irnprove concurrent access. For exarnple, if we have the relation on one disk and an index on another, accesses to the index can proceed without interfering with accesses to the relation, at least at the level of disk reads.

Replace Long Transactions by Short Ones: SometiInes, just too ruuch work is done within a transaction, and it takes a long tirne and holds locks a long tirne. Consider rewriting the transaction as two or Inore smaller transactions; holdable cursors (see Section 6.1.2) can be helpful in doing this. The advantage is that each new transaction cornpletes quicker and releases locks sooner. The disadvantage is that the original list of operations is no longer executed atomically, and the application code Illust deal with situations in which one or rnore of the new transactions fail.

Build a Warehouse: CC)lnplex queries can hold shared locks for a long tirne. ()ften, however, these queries involve statistical analysis of business trends and it is a,cceptable to run theln on a copy of the data that is a little out of date. rrhis led to the popularity of *data warehouses*, which are databases that cornplcnlcnt

the operational database by rnaintaining a copy of data used in cornplex queries (Chapter 25). Running these queries against the warehouse relieves the burden of long-running queries froln the operational database.

Consider a Lower Isolation Level: In rnany situations, such as queries generating aggregate infonnation or statistical sununaries, we can use a lower SQL isolation level such as REPEATABLE READ or READ COMMITTED (Section 16.6). Lo\ver isolation levels incur lower locking overheads, a,nel the application progrannner rllust make good design trade-offs.

20.10.2 Reducing Hot Spots

Delay Operations on Hot Spots: We already discussed the value of delaying lock requests. Obviously, this is especially irnportant for requests involving frequently used objects.

Optimize Access Patterns: The *patteTn* of updates to a relation can also be significant. For exanlple, if tuples are inserted into the Ernployees relation in *eid* order and we have a B+ tree index on *eid*, each insert goes to the last leaf page of the B+ tree. This leads to hot spots along the path froIn the root to the rightrnost leaf page. Such considerations nlay lead us to choose a hash index over a B+- tree index or to index on a different field. Note that this pattern of access leads to poor perforrnance for ISAM indexes as well, since the last leaf page beCOlIles a hot spot. rrhis is not a problcln for hash indexes because the hashing process randornizes the bucket into which a record is inserted.

Partition Operations on Hot Spots: Consider a data entry transaction that appends new records to a file (e.g., inserts into a table stored as a heap file). Instead of appending records one-per-transaction and obtaining a lock on the last page for each record, we can replace the transaction by several other transactions, each of which writes records to a local file and periodically appends a batch of records to the rnain file. While we do rnore work overall, this reduces the lock contention on the last page of the original file.

As a further illustration of partitioning, suppose we track the nU111ber of records inserted in a counter. Instead of updating this counter once per record, the preceding approach results in updating several counters and periodically updating the main counter. rrhis idea can be adapted to many uses of counters, \vith varying degrees of effort. For exaInple, consider a counter that tracks the Ilurnber of reservations, with the rule that a new reservation is allowed only if the counter is below a, rnaxiullun value. We can replace this by three counters, each \vith one-third the origina11naxirIlurn threshold, and three transactions that use these counters rather than the original. We obtain greater concurrency, but

have to deal with the case where one of the counters is at the maximum value but SOHIe other counter can still be incrcrnented. Thus, the price of greater concurrency is increased cornplexity in the logic of the application code.

Choice of Index: If a relation is updated frequently, B+ tree indexes can becolne a concurrency control bottleneck, because all accesses through the index HUlst go through the root. Thus, the root and index pages just below it can bec()lne hot spots. If the DBMS uses specialized locking protocols for tree indexes, and in particular, sets finc-granularity locks, this problenl is greatly alleviated. l\Ilany current systeuls use such techniques.

Nonetheless, this consideration lllay lead us to choose an ISAM index in SOlllle situations. Because the index levels of an ISAM index are static, we need not obtain locks on these pages; only the leaf pages need to be locked. An ISAl\!l index rnay be preferable to a B+ tree index, for exalllple, if frequent updates occur but we expect the relative distribution of records and the nUlnber (and size) of records with a given range of search key values to stay approxirnately the saIne. In this case the ISAM index offers a lower locking overhead (and reduced contention for locks), and the distribution of records is such that few overflow pages are created.

I-Iashed indexes do not create such a concurrency bottleneck, unless the data distribution is very skewed and lnany data itenlS are concentrated in a few buckets. In this case, the directory entries for these buckets can beccnne a hot spot.

20.11 CASE STUDY: THE INTERNET SHOP

Revisiting our running case study, I)BDudes considers the expected workload for the B&N 1)00kstore. rrhe owner of the bookstore expects rnost of his CIIS-torners to search for books by ISBN nUluber before placing an order. Placing an order involves inserting one record into the ()rders table and inserting one or lllore records into the Orderlists relation. If a sufficient nurnber of books is available, a, shiprnent is prepared and a value for the *ship_date* in the Orderlists relation is set. In addition, the available quantities of books in stock changes all the tirne, since orders are placed that; decrease the quantity available and new books arrive frorn suppliers and increase the quantity available.

The DBDudes tearn begins by considering searches for books by ISBN'. Since *isbn* is a key, (l,n equality query on *isbn* returns at rnost one record. rrhereforc, to speed up queries frolll Cllstolllers who look for books with a given ISBN, I)BIJudes decides to build an unclustered hash index on *isbn*.

Next, it considers updates to book quantities. To update the *qty_in_stock* value for a book, we IIIUSt first search for the book by ISBN; the index on *isbn* speeds this up. Since the *qty_in_stock* value for a book is updated quite frequently, DBDudes also considers partitioning the Books relation vertically into the fo1-1c)\ving two relations:

BooksQty(*isbn*, *qty*)
BookRest(*isbn*, *title*, *author*, *price*, *year_published*)

Unfortunately, this vertical partitioning slows do\vn another very popular query: Equality search on ISBN to retrieve all infonnation about a book now requires a join between BooksQty and BooksH,est. So DBDudes decides not to vertically partition Books.

DBDudcs thinks it is likely that custonlers will also want to search for books by title and by author, and decides to add unclustered hash indexes on *title* and author-these indexes are inexpensive to rnaintain because the set of books is rarely changed even though the quantity in stock for a book changes often.

Next, DBDudes considers the Custorners relation. A custorner is first identified by the unique custorner identifaction nurnber. So the rnost COlnrnon queries on Custorners are equality queries involving the custolner identification nurnber, and DBDudes decides to build a clustered hash index on *cid* to achieve maxirnum speed for this query.

l\!Ioving on to the Orders relation, DBDudes sees that it is involved in two queries: insertion of new orders and retrieval of existing orders. Both queries involve the *ordernum* attribute as search key and so DBDudes decides to huild an index on it. What type of index should this be—a 13+ tree or a hash index? Since order nurnbers are assigned sequentially and correspond to the order date, sorting by *ordernum* effectively sorts by order date as well. So DBDudes decides to build a clustered B+ tree index on *ordernum*. A.lthough the operational requirernents rncntioned until no\v favor neither a 13+ tree nor a hash index, B&N\vill probably want to rnonitor daily a,ctivities and the clustered 13+ tree is a better choice for such range queries. ()f course, this 1118ans that retrieving all orders for a given custorner could be expensive for custolllers with InallY orders, since clustering by *ordernum* precludes clustering by other attributes, SllCh as *cicio*

The ()rderlists rela,tion involves lnostly insertions, with an occasional update of a shiprnent date or a query to list all cOlnponents of a given order. If Orderlists is kept sorted on *ordernum*, all insertions are appends at the end of the relation and thus very efficient. A clustered 13+ tree index on *ordernum* maintains this sort order and also speeds up retrieval of aU iterns for a given order. To update

a shiprnent date, we need to search for a tuple by *ordernum* and *isbn.* The index on *ordernum* helps here as well. Although an index on ⟨*ordernum, isbn*⟩ would be better for this purpose, insertions would not be as efficient as with an index on just *ordernum*; DBDudes therefore decides to index ()rderlists on just *ordernum.*

20.11.1 Tuning the Database

Several rnonths after the launch of the B&N site, DBDudes is called in and told that custorner enquiries about pending orders are being processed very slowly. B&N has becorne very successful, and the Orders and Orderlists tables have grown huge.

l'hinking further about the design, DBDudes realizes that there are two types of orders: *completed orders,* for which all books have already shipped, and *partially co'mpleted orders*, for which sorne books are yet to be shipped. l\Ilost custorIler requests to look up an order involve partially corIlpleted orders, which are a sInall fraction of all orders. DBDudes therefore decides to horizontally partition both the Orders table and the Orderlists table by *ordernu'Tn.* This results in four new relations: NewOrders, OldOrders, NewOrderlists, and OldOrderlists.

An order and its cornponents are always in exactly one pair of relations····,·--and we can deterrIline which pair, old or new, by a sinlple check on *ordernum*—and queries involving that order can always be evaluated using only the relevant relations. SCHIle queries are now slower, such as those asking for all of a custoruer's orders, since they require us to search two sets of relations. lIowever, these queries are infrequent and their perforrnance is acceptable.

20.12 DBMS BENCHMARKING

Thus far, we considered how to irnprove the design of a database to obtain better perforrnance. 1\S the database grows, however; the underlying DBMS rnay no longer be able to provide adequate perforrnance, even with the best possible design, and we have to consider upgrading our systcrn, typically by buying faster harchva,re and additional rnernory. We IIlay also consider rnigrating our database to (1, new DBIVIS.

When evaluating DBMS products, perforrnal1ce is an iUlportant consideration. ADBIVIS is a cornplex piece of software, and different vendors rnay target their systerns toward different market segrnents by putting rnore effort into optirnizirlg certa,in parts of the systern or choosing different systern designs. For example, sorne systcrIls are designed to run cornplex queries efficiently, while others are designed to run Inany sirnple transactions per second. Within

each category of systcrIls, there are many cornpeting products. To assist users in choosing a DBMS that is 'well suited to their needs, several performance benchmarks have been developed. These include benchrnarks for measuring the perforlnance of a certain class of applications (e.g., the TPC benclnnarks) and benchrnarks for measuring how well a DBIVlS perfOrII1S various operations (e.g., the \Visconsin benchrnark).

Benchmarks should be portable, easy to understand, and scale naturally to larger problenl instances. They should II1eaSUre *peak performance* (e.g., *transactions per second,* or *ips)* as well as *pTice/perforrnance ratios* (e.g., *$/tps)* for typical workloads in a given application donlain. The Transaction Processing Council (TPC) was created to define benchlnarks for transaction processing and database systerns. Other well-known benchlnarks have been proposed by acadelnic researchers and industry organizations. Benchrnarks that are proprietary to a given vendor are not very useful for cornparing different systerns (although they rnay be useful in deterrnining how well a given systern would handle a particular workload).

20.12.1 Well-Known DBMS Benchmarks

Online Transaction Processing Benchmarks: The TPC-A and TPC-B benchrnarks constitute the standard definitions of the *ips* and $/*tps* measures. TPC-A rneasures the perfonnance and price of a computer network in addition to the DBMS, whereas the TPC-B benclnnark considers the DBMS by itself. These bencln11arks involve a sirnple transaction that updates three data records, frolll three different tables, and appends a record to a fourth table. A 11urnber of details (e.g., transaction arrival distribution, interconnect rnethod, systern properties) are rigorously specified, ensuring that results for different systenls can be rneaningfully cOI11pared. The T'PC-C benchmark is a l110re cornplex suite of transactional tasks than TPC-A and TPC-B. It rnodels a warehouse that tracks iterns supplied to custorners and involves five types of transactions. Each TPC-C transaction is rnuch rIlore expensive than a 1'PC-A or TPC-B transaction, a,nel TPC-C exercises a rnuch ,videI' range of systern capabilities, such as use of secondary indexes and transaction aborts. It has Inore or less cOlnpletely replaced TPC-A and rrpC-B as the standard transaction processing bencillnark.

Query Benchmarks: The Wisconsin l)cnchrnark is \videly used for 1neasnring the perforrnance of sirnple relational queries. The Set Query benclunark Hleasures the perforrnance of a suite of rJlore cornplex queries, and the $AS^3A.P$ l)enchrnark measures the perfonnance of (1, Inixed workload of transactions, relational queries, and utility fUllctions. The rrpC-I) benchn.lark is a suite of cornplex SQL queries intended to be representative of the (Incision-support ap-

plication domain. The ()LAP Council also developed a benchmark for cornplex decision-support queries, including some queries that cannot be expressed easily in SQL; this is intended to rneasure systerIls for *online a'nalyt'ic processing (OLAP)*,\vhieh we discuss in Chapter 25, rather than traditional SQL systerns. The Sequoia 2000 benchrnark is designed to cornpare DBMS support for geographic inforrnation systerns.

Object-Database Benchmarks: The 001 and 007 benclunarks rneasure the perforrnance of object-oriented database systelns. 'rhe Bucky benclunark rneasures the perforrnance of object-relational database systcrns. (We discuss object-database systelns in Chapter 23.)

20.12.2 Using a **Benchmark**

Benchrnarks should be used with a good understanding of what they are designed to rnea8ure and the application environrnent in \vhich a DBMS is to be used. When you use benchrnarks to guide your choice of a DBMS, keep the following guidelines in rnind:

- How Meaningful is a Given Benchmark? Benchrnarks that try to distill perforrnance into a single nunlber can be overly sirnplistic. A DBMS is a cOlnplex piece of software used in a variety of applications. A good benchlnark should have a suite of tasks that are carefully chosen to cover a particular application dornain and test DBMS features irnportant for that d01nain.

- How Well Does a Benchrnark Reflect Your Workload? Consider your expected workload and corupare it with the benchrnark. Give 11101'8 weight to the perfonnance of those l)enchrnark tasks (i.e., queries and updates) that are similar to irnportant tasks in your workload. Also consider how benclunark nurnbers are measured. For exarnple, elapsed tirne for individual queries rnight be rnisleading if considered in a rnultiuser setting: A systern may have higher elapsed times because of slower *l/C).* On a 1nultiuser workloa,d, given sufficient disks for parallel *l/C),* such a systern lnight olltperfofrn a SY8t8111 'with a lower elapsed time.

- Create Your Own Benchmark: Vendors often tweak their systerns in ad hoc ways to obtain good nurnbers on important benchmarks. fr0 counter this, create your own benc1unark by modifying standard benchrnarks slightly or by replacing the tasks in a standard benchrnark \vith similar tasks frarn your workload.

20.13 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are the cornponents of a workload description? (Section 20.1.1)
- What decisions need to be rnade during physical design? (Section 20.1.2)
- Describe six high-level guidelines for index selection. (Section 20.2)
- When should we create clustered indexes? (Section 20.4)
- What is co-clustering, and when should we use it? (Section 20.4.1)
- What is an index-only plan, and how do we create indexes for index-only plans? (Section 20.5)
- Why is automatic index tuning a hard problern? Give an exarnple. (Section 20.6.1)
- Give an exarnple of one algorithrn for autonlatic index tuning. (Section 20.6.2)
- Why is database tuning irnportant? (Section 20.7)
- How do we tune indexes, the conceptual scheula, and queries and views? (Sections 20.7.1 to 20.7.3)
- What are our choices in tuning the conceptual scherna? What are the following techniques and when should we apply thern: settling for a weaker norrnal forrn, denorrnalization, and horizontal and vertiacal decornpositions. (Section 20.8)
- What choices do we have in tuning queries and views? (Section 20.9)
- What is the irnpact of locking 011 database perforluance? I-Iow can we reduce lock contention and hot spots? (Section 20.10)
- Why do we have standaTdized database benclllnarks, and what conunon Inetrics are used to evaluate database systelns? Can you describe a few popular database benchrnarks? (Section 20.12)

EXERCISES

Exercise 20.1 Consider the following BCNF schcrna for a portion of a sirnple corporate database (type infonnation is not relevant to this question and is ornitted):

Ernp (*eid, ename, addr, sal, age, yrs, deptid)*
Dept (*did, dname, flooT, budget*)

Suppose you know that the following queries are the six most COUUllcm queries in the \vorkload for this corporation and that all six are roughly equivalent in frequency and inlportance:

- List the id, name, and address of eUlployees in a user-specified age range.
- List the id, name, and address of crnployees who work in the department \vith a llser-specified department narne.
- List the id and address of elnployees with a user-specified eluployeenanle.
- List the overall average salary for ernployees.
- List the average salary for eInployees of each age; that is, for each age in the datal)(1se, list the age and the corresponding average salary.
- List all the departrnent infonnation, ordered by departrnent floor nurnbers.

1. Given this infonnation, and assuIning that these queries are lnore iluportant than any updates, design a physical scherna for the corporate database that will give good perfor-rnance for the expected workload. In particular, decide which attributes will be indexed and whether each index will be a clustered index or an unclustered index. Assume that 13+ tree indexes are the only index type supported by the DBMS and that both single- and nnrltiple-attribute keys are pernlitted. Specify yOllr physical design by identifying the attributes you recornnlCnd indexing on via clustered or unclustered 13+ trees.
2. Redesign the physical schelna assuIning that the set of iInportant queries is changed to be the following:
   - List the id and address of enlployees with a user-specified ernployee narne.
   - List the overall rnaxinHun salary for eruployees.
   - List the average salary for ernployees by departnlent; that is, for each *deptid* value, list the *deptid* value and the average salary of ernployees in that departrnent.
   - List the Slun of the budgets of all departrnents by floor; that is, for each floor, list the floor and the sum.
   - AssuIne that this workload is to be tuned with an autornatic index tuning wizard. Outline the rnain steps in the execution of the index tuning algorithrn and the set of candidate configurations that would be considered.

Exercise 20.2 Consider the follo\ving BCNF' relational scherna for a portion of a universit.y database (type infonnation is not relevant to this question and is ornitted):

Prof(*ssno*, *pname*, *office*, *age*, *sex*, *specialty*, *dept_did*)
Dept(*did*, *dname*, *budget*, *num_majors*, *chair_ssno*)

Suppose you kno\v that the following queries are the five rnost connnon queries in the workloa,d for this university and that all five are roughly equivalent in frequency and importance:

- List the names, ages, and offices of professors of a user-specified sex (rnale or fernale) who have a user-specified research specialty (e.g., *recursive query processing*). Assurne that the university has a diverse set of faculty members, rnaking it very unCOlnmon for Inore than a fe\\! professors to have the same research specialty.
- List all the departrnent information for departrnents with professors in a user-specified age range.
- List the <lepartlnent i<l, department name, and chairperson name for departments with a user-specified nurnber of majors.

- List the lowest budget for a department in the university.
- List all the infornultion about professors \vho are departlnent chairpersons.

These queries occur runch 1nore frequently than updates, so you should build whatever indexes you need to speed up these queries. However, you should not build any unnecessary indexes, as updates will occur (and would be slowed down by unnecessary indexes). Given this information, design a physical schc1na for the university database that will give good perfonnance for the expected workload. In particular, decide which attributes should be indexed and 'whether each index should be a clustered index or an unclustered index. Assulne that both B+ trees and hashed indexes are supported by the DBMS and that both single- and lImltiple-attribute index search keys are perrnitted.

1. Specify your physical design by identifying the attributes you recomlnend indexing on, indicating whether each index should be clustered or unclustered and whether it should be a B+ tree or a hashed index.
2. Assurne that this workload is to be tuned with an autornatic index tuning wizard. Outline the rnain steps in the algorithrn and the set of candidate configurations considered.
3. Redesign the physical schema, assurning that the set of irnportant queries is changed to be the following:
   - List the nUlnber of different specialties covered by professors in each department, by department.
   - Find the departrnent with the fewest rnajors.
   - Find the youngest professor who is a department chairperson.

Exercise 20.3 Consider the following BCNF relational schmna for a portion of a cornpany database (type inforrnation is not relevant to this question and is OInitted):

Project*(pno, proj_name, proj_base_dept, proj_mgr, topic, budget)*
Manager(*mid, mgr_name, mgr_dept, salary, age, sex:)*

Note that each project is based in sorne department, each manager is e1Ylployed in some departIllEmt, and the manager of a project need not be e1nployed in the sarne departrnent (in which the project is based). Suppose you know that the following queries are the five most COHUllon queries in the workload for this university and all five are roughly equivalent in frequency and i1"nportance:

- List the names, ages, and salaries of lnanagers of a user-specified sex (rnale or female) working in a given department. You can assurne that, while there are rnany departments, each department contains very few project managers.
- List the names of all projects with lnanagers whose ages are m a user-specified range (e.g., younger than 30).
- List the names of all departrnents such that a rnanager III this department manages a project based in this department.
- List the name of the project with the lowest budget.
- List the names of all managers in the SaITIC department as a given project.

These queries occur nIllch more frequently than updates, so you should build \vhatever indexes you need to speed up these queries. However, you should not build any unnecessary indexes, as updates \'lill occur (a"nd \vould be slowed down by urmccessary indexes). Given

this infonnatioll, design a physical schema for the conlpany database that win give good performance for the expected \vorkload. In particular, decide which attributes should be indexed and whether each index should be a clustered index or an unclustered index. Assuille that both B+ trees and hashed indexes are supported by the DBMS, and that both single- and Illuitiple-attribute index keys are perruitted.

1. Specify your physical design by identifying the attributes you recOlIlrIlend indexing on, indicating whether each index should be clustered or unclustered and whether it should be a B+ tree or a hashed index.
2. Assunle that this workload is to be tuned with an autornatic index tuning wizard. Outline the lnain steps in the algorithrn and the set of candidate configurations considered.
3. Redesign the physical schema assulning the set of ilnportant queries is changed to be the following:
   - Find the total of the budgets for projects luanaged by each rnanager; that is, list *p'roj_rngr* and the total of the budgets of projects luanaged by that manager, for all values of *proj_mgT.*
   - Find the total of the budgets for projects managed by each rnanager but only for managers who are in a user-specified age range.
   - Find the number of male rnanagers.
   - Find the average age of rnanagers.

Exercise 20.4 The Globetrotters Club is organized into chapters. The president of a chapter can never serve as the president of any other chapter, and each chapter gives its president sonle salary. Chapters keep moving to new locations, and a new president is elected when (and only when) a chapter rnoves. This data is stored in a relation *G(C,S,L,P),* where the attributes are chapters *(C),* salaries *(S)*, locations *(L)*, and presidents *(P).* Queries of the following fornl are frequently asked, and you *mU8t* be able to answer thern without cOluputing a join: "Who was the president of chapter X when it was in location *Y*?"

1. List the FDs that are given to hold over G.
2. What are the candidate keys for relation G?
3. What Honnal fornl is the scherna Gin?
4. Design a good database scherna for the club. (Rernernber that your design *must* satisfy the stated query requirenlent!)
5. What nonnal fonn is your good scherna in? Give an exarnple of a query that is likely to run slower on this schema than on the relation G.
6. Is there a lossless-join, dependency-preserving deCOlTlposition of G into BeNF?
7. Is there ever a good reason to accept sornething less than :3NF' \vhen designing a schema for a relational database? Use this example, if necessary adding further constraints, to illustrate your answer.

Exercise 20.5 Consider the following BCNF relation, which lists the ids, types (e.g., nuts or bolts), and costs of various parts, along with the mllnber available or in stock:

Parts (*pid, pname, cost, num_avail*)

You are told that the following two queries are extrelnely irnportant:

- Find the total nunlber available by part type, for all types. (That is, the surn of the *num_avail* value of all nuts, the sum of the *num_avail* value of all bolts, and so forth)
- List the *pids* of parts with the highest cost.

1. Describe the physical design that you would choose for this relation. That is, what kind of a file structure would you choose for the set of Parts records, and what indexes would you create?
2. Suppose your custorners subsequently cmnplain that performance is still not satisfactory (given the indexes and file organization you chose for the Parts relation in response to the previous question). Since you cannot afford to buy new hardware or software, you have to consider a schenla redesign. Explain how you would try to obtain better perfonnance by describing the scherna for the relation(s) that you would use and your choice of file organizations and indexes on these relations.
3. How would your answers to the two questions change, if at all, if your systeIl1 did not support indexes with multiple-attribute search keys?

Exercise 20.6 Consider the following BCNF relations, which describe ernployees and the departments they work in:

Ernp (*eid, sal, did)*
Dept (*d'id, location, budget*)

You are told that the following queries are extrernely important:

- Find the location where a user-specified enlployee works.
- Check whether the budget of a department is greater than the salary of each ernployee in that departrnent.

1. Describe the physical design you would choose for this relation. That is, what kind of a file structure would you choose for these relations, and what indexes would you create?
2. Suppose that your custollwrs subsequently cOIuplain that perforrnance is still not satisfactory (given the indexes and file organization that you chose for the relations in response to the previous question). Since you cannot afford to buy new hardware or software, you have to consider a schelna redesign. Explain how you would try to obtain better perfonnance by describing the scherna for the relation(s) that you would use and your choice of file organizations and indexes on these relations.
3. Suppose that your database systern has very ineff1cient irnplenlentations of index structures. What kind of a design would you try in this case?

Exercise 20.7 Consider the following BCNF relations, which describe departrnents in a company and ernployees:

Dept(*did, dname, location managerid*)
Enlp(*eid, sal*)

'You are told that the followving queries are extrernely iruportant:

- List the names and ids of rnanagel's for each department in a user-specified location., in alphabetical order by department narne.
- Find the average salary of ernployees who manage departments in a user-specified location.You can assume that no one rnanages nlOre than one depa,rtrnent.

1. Describe the file structures and indexes that you would choose.
2. You subsequently realize that updates to these relations are frequent. Because indexes incur a high overhead, can you think of a way to irnprove perforrnance on these queries without using indexes?

Exercise 20.8 For each of the following queries, identify one possible reason why an opti-Inizer Illight not find a good plan. Rewrite the query so that a good plan is likely to be found. Any available indexes or known constraints are listed before each query; assurne that the relation schelnas are consistent with the attributes referred to in the query.

1. An index is available on the *age* attribute:

```
SELECT E.dno
FROM   Elnployee E
WHERE  E.age=20 OR E.age=10
```

2. A B+ tree index is available on the *age* attribute:

```
SELECT E.dno
FROM   Employee E
WHERE  E.age<20 AND E.age>10
```

3. An index is available on the *age* attribute:

```
SELECT E.eIno
FROM   Enlployee E
WHERE  2*E.age<20
```

4. No index is available:

```
SELECT DISTINCT *
FROM   Enlployee E
```

5. No index is available:

```
SELECT   AVG (B.sal)
FROM     Elnployee E
GROUP BY E.dno
HAVING   E.dno=22
```

6. The *sid* in Reserves is a foreign key that refers to Sailors:

```
SELECT   S.sid
FROM     Sailors S, Reserves H
WHERE    S.sid=R.sid
```

Exercise 20.9 Consider two 'ways to COlupute the names of elnployees who earn rnore than $100,000 and whose age is equal to their manager's age. First, a nested query:

```
SELECT   E1.ename
FROM     Emp E1
WHERE    E1.sal > 100 AND E1.age = ( SELECT E2.age
                                     FROM   Ernp E2, Dept D2
                                     WHERE  E1.dname = D2.dname
                                            AND D2.mgr = E2.enarne )
```

Second, a query that uses a view definition:

```
SELECT   £1.enarne
FROM     Ernp E1, MgrAge A
WHERE    El.dnarue = A.dnarne AND E1.sal > 100 AND E1.age = A.age

CREATE VIEW MgrAge (dnan1e, age)
       AS SELECT D.dnanw, E.age
          FROM   Errlp E, Dept D
          WHERE  D.mgr = E.ename
```

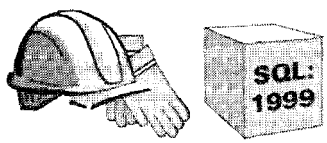
1. Describe a situation in which the first query is likely to outperforrn the second query.
2. Describe a situation in which the second query is likely to outperfonn the first query.
3. Can you construct an equivalent query that is likely to beat both these queries when every ernployee who earns rnore than $100,000 is either 35 or 40 years old? Explain briefly.

BIBLIOGRAPHIC NOTES

[658] is an early discussion of physical database design. [659] discusses the performance implications of normalization and observes that denormalization may improve perforrnance for certain queries. The ideas underlying a physical design tool frorn IBl'vf are described in [272]. The Microsoft AutoAdrnin tool that perfonns automatic index selection according to a query workload is described in several papers [163, 164]. The DB2 Advisor is described in [750]. Other approaches to physical database design are described in [146, 639]. [679] considers *transaction tuning*, which we discussed only briefly. The issue is how an application should be structured into a collection of transactions to rnaxirnize perfonnance.

The following books on database design cover physical design issues in detail; they are reCOIll-rnended for further reading. [274] is largely independent of specific products, although rnaBy examples are based on DB2 and Teradata systerl1S. [779] deals prirnarily with DB2. Shasha and Bonnet give an in-depth, readable introduction to database tuning [104].

[334] contains several papers on benchrnarking database systerns and has accompanying soft-ware. It includes articles on the AS^3AP, Set Query, 'I'PC-A, 'rpC-B, Wisconsin, and 001 bendunarks written by the original developers. The Bucky benchrnark is described in [132], the 007 benchrnark is described in [131], and the T'pe-D benchrnark is described in [739]. The Sequoia 2000 bendunark is described in [720].



21

SECURITY AND AUTHORIZATION

- What are the rnain security considerations in designing a database application?
- What IIlechanisms does a DBMS provide to control a user's access to data?
- What is discretionary access control and how is it supported in SQL?
- What are the weaknesses of discretionary access control? How are these addressed in lnandatory access control?
- What are covert channels and how do they cornpromise lnandatory access control?
- What lTIUst the DBA do to ensure security?
- What is the added security threat when a database is accessed re-rnotely?
- What is the role of encryption in ensuring secure access? How is it used for certifying servers and creating digital sig11atures?
- Key concepts: security, integrity, availability; discretionary access control, privileges, GRANT, REVOKE; rna.ndatory access control, objects, subjects, security classes, rnultilevel tables, polyinstantiation; covert channels, DoD security levels; statistical databases, inferring secure information; authentication for reIllote access, securing servers, digital signatures; encyption, public-key encryption.

I know that's a secret, for it's whispered everywhere.

William Congreve

The data stored in a DBMS is often vital to the business interests of the organization and is regarded as a corporate asset. In addition to protecting the intrinsic value of the data, corporations rnust consider O\vays to ensure privacy and control access to data that must not be revealed to certain groups of users for various reasons.

In this chapter, we discuss the concepts underlying access control and security in a DBMS. After introducing database security issues in Section 21.1,we consider two distinct approaches, called *discretionary* and *mandatory*, to specifying and lTlanaging access controls. An access control Inechanism is a way to control the data accessible by a given user. After introducing access controls in Section 21.2, we cover discretionary access control,which is supported in SQL, in Section 21.3.vVe briefly cover n1andatory access control, which is not supported in SQL, in Section 21.4.

In Section 21.6, we discuss SOIne additional aspects of database security, such as security in a statistical database and the role of the database adrninistrator. We then consider SOlne of the unique challenges in supporting secure access to a DBMS over the Internet, which is a central problern in e-COlllInerce and other Internet database applications, in Section 21.5. We conclude this chapter with a discussion of security aspects of the Barns and Nobble case study in Section 21.7.

21.1 INTRODUCTION TO DATABASE SECURITY

There are three rnain objectives when designing a secure database application:

1. Secrecy: InfoI'rnation should not be disclosed to unauthorized users. For exarnple, a student should not be allowed to exarnine other students' grades.
2. Integrity: ()nly authorized users should be allowed to Hlodify data. For example, students 1Ilay be allowed to see their grades, yet not allowed (obviously) to rnodify thern.
3. Availability: Authorized users should not be denied access. For example, an instructor who wishes to change a grade should be allowed to do so.

To achieve these objectives, a clear and consistent security policy should be developed to describe what security measures rnust be enforced. In particular, we rnu8t detennine what part of the data is to be protected and which users get access to which portions of the data. Next, the security mechanisrns of the underlying I)B:JVIS and operating systenl, as well as externaJ mechanisms,

such as securing access to buildings, Illust be utilized to enforce the policy. We emphasize that security rneasures IIlust l)e taken at several levels.

Security leaks in the OS or network connections can cirCUlnvent database security rnechanisrns. For example, such leaks could allow an intruder to log on as the database acbninistrator, 'with all the attendant I)BlVIS access rights. Hurnan factors are another source of security leaks. :For exarnple, a user IHay choose a password that is easy to guess, or a user who is authorized to see sensitive data may luisuse it. Such errors account for a large percentage of security breaches. \Ve do not discuss these aspects of security despite their irllportance because they are not specific to database rnanagerllent systelIls; our IIlain focus is on database access control rllechanisrns to support a security policy.

We observe that vie\vs are a valuable tool in enforcing security policies. The view rnechanisrll can be used to create a 'window' 011 a collection of data that is appropriate for SOllIe group of users. 'Views allow us to liUlit access to sensitive data by providing access to a restricted version (defined through a view) of that data, rather than to the data itself.

We use the following schemas in our exaurples:

Sailors(*sid:* integer, *snarne:* string, *rating:* integer, *age:* real)
Boats(*bid:* integer, *bnarne:* string, *color:* string)
Reserves(*sid:* integer, *bid:* integer, *day:* dates)

Increasingly, as database systcrlls becorne the backbone of e-COlluncrce applications requests originate over the Internet. This rnakes it irnportant to be able to authenticate a user to the database systern. A.fter all, enforcing a security policy that allows user Sarn to read a table and Ehner to write the table is not of l11uch use if Sam can rnasquerade as Ebner. COllversely, we Inus!; be able to assure users that they a,re COIIuIlunicating \vith a legitilnate systern (e.g., the real Arnazoll.col11 server, and not a spurious application intended to steal sensitive inforrnation such as a credit card nurl11>cr). vVhile the details of authentication are outside the scope of our coverage, we discuss the role of authentication and the l)Hsic ideas involved in Section 21.5, after covering database access control rnechanisrIls.

21.2 ACCESS CONTROL

A database for an enterprise contains a great deal of inforrnation and usually has several groups of users. 1\IJost users need to access onl,y a sruall pa;rt of the database to carry out their tasks. J\l1owing users unrestricted access to all the

data can be undesirable, and a DBMS should provide rnechanisHls to control access to data.

A DBMS offers two rnain approaches to access control. Discretionary access control is based on the concept of access rights, or privileges, and rnechanisrllS for giving users such privileges. A privilege allows a user to access Borne data object in a certain IIlHnIler (e.g., to read or 11lOdify). A user who creates a database object such as a table or a view autornatically gets all applicable privileges on that object. The D.BMS subsequently keeps track of how these privileges are granted to other users, and possibly revoked, and ensures that at all tirnes only users with the necessary privileges can access all object. SQL supports discretionary access control through the GRANT and REVOKE conunands. The GRANT cOllnnand gives privileges to users, and the REVOKE cornrnand takes away privileges. We discuss discretionary access control in Section 21.3.

Discretionary access control rnechanisrns, while generally effective, have certain weaknesses. In particular, a devious unauthorized user can trick an authorized user into disclosing sensitive data. Mandatory access control is based on systemwide policies that cannot be changed by individual users. In this approach each database object is assigned a *security class,* each user is assigned *clearance* for a security class, and rules are irnposed on reading and writing of database objects by users. The DBMS deterrnines whether a given user can read or write a given object based on certain rules that involve the security level of the object and the clearance of the user. These rules seek to ensure that sensitive data can never be 'passed on' to a user without the necessary clearance. 'rhe SQL standard does not include any support for rnandatory access control. 'We discuss rnandatory access control in Section 21.4.

21.3 DISCRETIONARY ACCESS CONTROL

SQL supports discretionary access control through the GRANT and REVOKE cornrnands. The GRANT cornrnand gives users privileges to base tables and views. 'rhe syntax of this command is as follows:

GRANT privileges ON object TO users [WITH GRANT OPTION]

For our purposes object is either a base table or a view. SQL recognizes certain other kinds of objects, but we do not discuss thcrn. Several privileges can be specified, including these:

- SELECT: The right to access (read) all colurnns of the table specified as the object, *including columns added later* through ALTER TABLE cornrnands.

- INSERT(*column-name*): The right to insert rows with (non-*null* or non-default) values in the naTned cohnnn of the table named as object. If this right is to be granted with respect to all cohunns, including coluulns that rnight be added later, we can sirnply usc INSERT. The privileges UPDATE(*column-name*) and UPDATE are sirnilar.
- DELETE: The right to delete rows frorn the table narned as object.
- REFERENCES(*col'Urnn-namJe):* The right to define foreign keys (in other tables) that refer to the specified cohnnn of the table object. REFERENCES without a colurnn naUIe specified denotes this right with respect to all colurnns, including any that are added later.

If a user has a privilege with the grant option, he or she can pass it to another user (with or without the grant option) by using the GRANT conunand. A user who creates a base table autolnatically has all applicable privileges on it, along with the right to grant these privileges to other users. A user who creates a view has precisely those privileges on the view that he or she has on *everyone* of the views or base tables used to define the view. The user creating the view Inust have the SELECT privilege on each underlying table, of course, and so is always granted the SELECT privilege on the view. The creator of the view has the SELECT privilege with the grant option only if he or she has the SELECT privilege with the grant option on every underlying table. In addition, if the view is updatable and the user holds INSERT, DELETE, or UPDATE privileges (with or without the grant option) on the (single) underlying table, the user autornatically gets the same privileges on the view.

()nly the owner of a scherna can execute the data definition statcrnents CREATE, ALTER, and DROP on that schcrna. The right to execute these staternents cannot be granted or revoked.

In conjullction with the GRANT and REVOKE cOl1llnands, views are an irnportant cornponent of the security rnechanisrns provided by a relational DBMS. By defining views on the base tables, we can present needed inforrnation to a user "while *hiding* other inforrnation that the user should not be given access to. For example, consider the following view definition:

```
CREATE VIEW }\ctiveSajlors (name, age, day)
        AS SELECT S.snarne, S.age, R.day
           FROM   Sailors S, Reserves R
           WHERE  S.sid = R.sid AND S.rating > 6
```

A user who can access ActiveSailors but not Sailors or Reserves kno\vs the names of sailors who have reservations but cannot find out the *bids* of boats reserved by a given sailor.

Role-Based Authorization in SQL: Privileges are assigned to users (authorization IDs, to be precise) in SQL-92. In the real world, privileges are often associated with a user's job or *role* within the organizat;ion. Many DBMSs have long supported the concept of a role and allowed privileges to be assigned to roles. Roles can then he granted to users and other roles. (Of courses, privileges can also be granted directly to users.) l'he SQL:1999 standard includes support for roles. Roles can be created and destroyed using the CREATE ROLE and DROP ROLE eornrnands. Users can be granted roles (optionally, with the ability to pass the role on to others). The standard GRANT and REVOKE connnands can assign privileges to (and revoke from) roles or authorization IDs.

What is the benefit of including a feature that Inany systerns already support? 'This ensures that, over tiIne, all vendors who comply with the standard support this feature. 'rhus, users can use the feature without worrying about portability of their application across DBMSs.

Privileges are assigned in SQL to authorization IDs, which can denote a single user or a group of users; a user lllUSt specify an authorization ID and, in Inany systerns, a corresponding *password* before the DBMS accepts any C0111-rnancls from hirn or her. So, technically, *Joe*, *Michael*, and so on are authorization IDs rather than user nan1es in the following exalllpies.

Suppose that user Joe has created the tables Boats, Reserves, and Sailors. Senne exarnples of the GRANT cOllunand that Joe can now execute fo11o\v:

```
GRANT INSERT, DELETE ON Reserves TO Yuppy WITH GRANT OPTION
GRANT SELECT ON Reserves TO Michael
GRANT SELECT ON Sailors TO Michael WITH GRANT OPTION
GRANT UPDATE (rating) ON Sailors TO Leah
GRANT REFERENCES (bid) ON Boats TO Bill
```

Yuppy can insert or delete Reserves rows and authorize SOlneone else to do the sarne. I\1ichael can execute SELECT queries on Sailors and H,eserves, and 118 can pass this privilege to others for Sailors but not for R,eserves. With the SELECT privilege, Michael can create a view that accesses the Sailors and Reserves tables (for example, the ActiveSailors vic\v), but he cannot grant SELECT on ActiveSailors to others.

()rI the other hand, suppose that Michael creates the foUo\ving view:

```
CREATE VIEWYoungSailors (sicl, age, rating)
        AS SELECT S.sicl, S.age, S.rating
```

```
FROM    Sailors S
WHERE   S.age < 18
```

The only underlying table is Sailors, for which Michael has SELECT with the grant option. He therefore has SELECT with the grant option on YoungSailors and can pass on the SELECT privilege on YoungSailors to Eric and Guppy:

```
GRANT SELECT ON YoungSailors TO Eric, Guppy
```

Eric and Guppy can now execute SELECT queries on the view YoungSailors-note, however, that Eric and Guppy do *not* have the right to execute SELECT queries directly on the underlying Sailors table.

Michael can also define constraints based on the inforrnation in the Sailors and Reserves tables. For exarnple, Michael can define the following table, which has an associated table constraint:

```
CREATE TABLE Sneaky (lnaxrating   INTEGER,
                     CHECK (maxrating >=
                            ( SELECT MAX (S.rating)
                              FROM    Sailors S )))
```

By repeatedly inserting rows with gradually increasing *rnaxrating* values into the Sneaky table until an insertion finally succeeds, lVIichael can find out the highest *rating* value in the Sailors table. This exarnple illustrates why SQL requires the creator of a table constraint that refers to Sailors to possess the SELECT privilege on Sailors.

Returning to the privileges granted by Joe, Leah can update only the *rating* colulnn of Sailors rows. She can execute the following cornmand, which sets all ratings to 8:

```
UPDATE  Sailors S
SET     S.rating = 8
```

However, she cannot execute the same cOllunand if the SET clause is changed to be SET *S.age* = 25, because she is not allowed to update the *age* field. A rnoro subtle point is illustrated by the following cOIrllnand, which decrelnents the rating of all 'sailors:

```
UPDATE  Sailors S
SET     S.ratillg = S.rating-l
```

Leah cannot execute this cOlInnand because it requires the SELECT privilege 011 the *S.rating* colurnn anei Leah does not have this privilege.

Bill can refer to the *bid* colurnn of Boats as a foreign key in another table. For example, Bill can create the Reserves table through the following cOlnnland:

```
CREATE TABLE R"eserves (sid     INTEGER,
                        bid     INTEGER,
                        day     DATE,
                        PRIMARY KEY (bid, day),
                        FOREIGN KEY (sid) REFERENCES Sailors ),
                        FOREIGN KEY (bid) REFERENCES Boats)
```

If Bill did not have the REFERENCES privilege on the *bid* coh1111n of Boats, he would not be able to execute this CREATE staternent because the FOREIGN KEY clause requires this privilege. (A sirnilar point holds with respect to the foreign key reference to Sailors.)

Specifying just the INSERT privilege (sirnilarly, REFERENCES and other privileges) in a GRANT conlmand is not the sarne as specifying SELECT(*colurnn-name)* for each column currently in the table. Consider the following command over the Sailors table, which has cohllnns *sid*, *snarne, rating,* and *age:*

```
GRANT INSERT ON Sailors TO J\!Iichael
```

Suppose that this conunand is executed and then a colurnn is added to the Sailors table (by executing an ALTER TABLE cOlIllnand). Note that Michael has the INSERT privilege with respect to the newly added colurnn. If we had executed the following GRANT cornrnand, instead of the previous one, Michael would not have the INSERT privilege on the new cohllnn:

```
GRANT   INSERT ON Sailors(sid), Sailors(sna1ne) , Sailors(rating),
        Sailors( age), TO J\!Iichael
```

There is a cornplernentary corl1rnand to GRANT that allows the withdrawal of privileges. The syntax of the REVOKE cOllunand is as follows:

```
REVOKE [GRANT OPTION FOR ] privileges
       ON object FROM users {RESTRICT | CASCADE }
```

The cOIlnnand CH,n be used to revoke either a privilege or just the grant option on a privilege (by using the optional GRANT OPTION FOR clause). One of the two alternatives, RESTRICT or CASCADE, HUlst be specified; we see 'what this choice IneaI1S shortly.

The intuition behind the GRANT command is clear: rrhe creator of a base table or a view is given all the appropriate privileges \vith respect to it and is allowed

to pass these privileges—including the right to pass along a privilege—to other users. The REVOKE comuland is, as expected, intended to achieve the reverse: A user who has granted a privilege to another user rnay change his or her lnincI and want to withdraw the gra,nted privilege. The intuition behind exactly 'what effect a REVOKE cornrnand has is conlplicated by the fact that a user Inay be granted the sarne privilege rnultiple tilnes, possibly by different users.

\Vhen a user executes a REVOKE cornmand with the CASCADE keyword, the effect is to \vithdraw the named privileges or grant option froIn all users who currently hold these privileges *solely* through a GRANT cOllunand that was previously executed by the sallIe user who is now executing the REVOKE cOl1nnand. If these users received the privileges with the grant option and passed it along, those recipients in turn lose their privileges as a consequence of the REVOKE cOlurnand, unless they received these privileges through an additional GRANT comIuand.

We illustrate the REVOKE cOllllnand through several examples. First, consider what happens after the following sequence of eornmands, where Joe is the creator of Sailors.

```
GRANT SELECT ON Sailors TO Art WITH GRANT OPTION   (executed by Joe)
GRANT SELECT ON Sailors TO Bob WITH GRANT OPTION   (executed by Art)
REVOKE SELECT ON Sailors FROM Art CASCADE          (executed by Joe)
```

Art loses the SELECT privilege on Sailors, of course. Then Bob, who received this privilege from Art, and only Art, also loses this privilege. Bob's privilege is said to be abandoned when the privilege froIn which it was derived (Art's SELECT privilege with grant option, in this exarnple) is revoked. When the CASCADE keyword is specified, all abandoned privileges are also revoked (possibly causing privileges held by other users to becOlne abandoned and thereby revoked recursively). If the RESTRICT keyword is specified in the REVOKE corllmand, the cornrnand is rejected if revoking the privileges *just* frorn the users specified in the cOIllluand would result in other privileges becorning abandoned.

Consider the following sequence, as another exarnple:

```
GRANT SELECT ON Sailors TO Art WITH GRANT OPTION   (executed by Joe)
GRANT SELECT ON Sailors TO Bob WITH GRANT OPTION   (executed by Joe)
GRANT SELECT ON Sailors TO Bob WITH GRANT OPTION   (executed by Art)
REVOKE SELECT ON Sailors FROM Art CASCADE          (executed by Joe)
```

As before, Art loses the SELECT privilege on Sailors. But what about Bob? Bob received this privilege fronl Art, but he also received it independently

(coincidentally, directly fro111 Joe). So Bob retains this privilege. Consider a third example:

```
GRANT SELECT ON Sailors TO Art WITH GRANT OPTION   (executed by Joe)
GRANT SELECT ON Sailors TO Art WITH GRANT OPTION   (executed by Joe)
REVOKE SELECT ON Sailors FROM Art CASCADE          (executed by Joe)
```

Since Joe granted the privilege to Art twice and only revoked it once, does Art get to keep the privilege? As per the SQL standard, no. Even if Joe absentmindedly granted the saIne privilege to Art several tirnes, he can revoke it with a single REVOKE cOIlunand.

It is possible to revoke just the grant option on a, privilege:

```
GRANT SELECT ON Sailors TO Art WITH GRANT OPTION   (executed by Joe)
REVOKE GRANT OPTION FOR SELECT ON Sailors
      FROM Art CASCADE                             (executed by Joe)
```

This cOlnmand would leave Art with the SELECT privilege on Sailors, but Art no longer has the grant option on this privilege and therefore cannot pass it on to other users.

These exarnples bring out the intuition behind the REVOKE cOilllland, and they highlight the cOlllplex interaction between GRANT and REVOKE cOlnnlands. When a GRANT is executed, a privilege descriptor is added to a table of such descriptors Inaintained by the DElVIS. The privilege descriptor specifies the following: the *grantor* of the privilege, the *grantee* who receives the privilege, the *granted privilege* (including the narne of the object involved), and whether the grant option is included. When a user creates a table or view and 'autornatically' gets certain privileges, a privilege descriptor with *system,* as the grantor is entered into this table.

rrhe effect of a series of GRANT cornrnands can be described in terrns of an authorization graph in which the nodes are users---technically, they are authorization IDs---and the arcs indicate .how privileges are passed. There is an arc fron1 (the node for) user 1. to user 2 if user 1. executed a GRANT cOIIunand giving a privilege to user 2; the arc is labeled with the descriptor for the GRANT cOIlllnand. A GRANT cOIlnnand has no effect if the saIne privileges ha.ve already been granted to the same grantee by th.e sarne grantor. The following sequence of commands illustrates the sernantics of GRANT and REVOKE connnands when there is a *cycle* in the authorization graph:

```
GRANT SELECT ON Sailors TO .Art WITH GRANT OPTION   (executed by Joe)
GRANT SELECT ON Sa.ilors TO Bob WITH GRANT OPTION   (executed by Art)
```

```
GRANT SELECT ON Sailors TO Art WITH GRANT OPTION   (executed by Bob)
GRANT SELECT ON Sailors TO Cal WITH GRANT OPTION   (executed by Joe)
GRANT SELECT ON Sailors TO Bob WITH GRANT OPTION   (executed by Cal)
REVOKE SELECT ON Sailors FROM Art CASCADE          (executed by Joe)
```

The authorization graph for this exarnple is shown in Figure 21.1. Note that we indicate how Joe, the creator of Sailors, acquired the SELECT privilege frorl1 the DBMS by introdtIcing a *System* node and drawing an arc froIn this node to Joe's node.

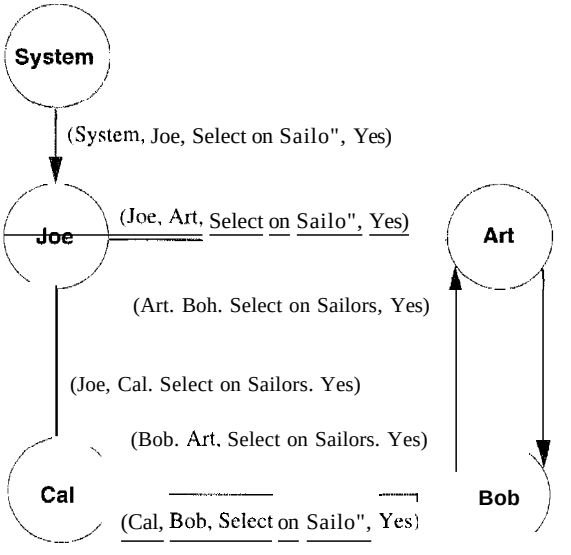


Figure 21.1 Example Authorization Graph

As the graph dearly indicates, Bob's grant to Art and Art's grant to Bob (of the same privilege) creates a cycle. Bob is subsequently given the saIne privilege by Cal, who received it independently froIn Joe. At this point Joe decides to revoke the privilege he granted Art.

Let us trace the effect of this revocation. The arc [raIn Joe to Art is removed because it corresponds to the granting action that is revoked. All rernaining nodes have the following property: *1f node N has an outgoing arc labeled with a privilege, there is a path fTorn the System node to 'node N in 'which each aTC label contains the same privilege plus the grant opt'ion.* That is, any rernaining granting action is justified by a privilege received (directly or indirectly) frorn the Systern. The execution of Joe's REVOKE conllnand therefore stops at this POiIlt,\vith everyone continuing to hold the SELECT privilege on Sailors.

rrhis result may seenl nnintuitive because Art continues to have the privilege only because he received it fr0111 Bob, and at the time that Bob granted the privilege to Art, he had received it only frorn Art. Although Bob acquired the privilege through Cal subsequently, should we not undo the effect of his grant

to Art when executing Joe's REVOKE command? 'rhe effect of the grant frorn Bob to Art is *not* undone in SQL. In effect. if a user acquires a privilege rnultiple tilnes frolIl different grantors, SQL treats each of these grants to the user as having occurred *befoTe* that user passed on the privilege to other users. This implementation of REVOKE is convenient in 111any reaJ-\vorld situations. .For exanlple, if a lIl.anager is fired after passing on sorne privileges to subordinates (who lnay in turn have passed the privileges to others), we can ensure that only the rnanager's privileges are rernoved by first redoing all of the Illanager's granting actions and then revoking his or her privileges. That is, we need not recursively redo the subordinates' granting actions.

To return to the saga of Joe and his friends, let us suppose that Joe decides to revoke Cal's SELECT privilege as well. Clearly, the arc frorn Joe to Cal corresponding to the grant of this privilege is rerIloved. The arc frorH Cal to Bob is reilloved as well, since there is no longer a path fronl SystelIl to Cal that gives Cal the right to pass the SELECT privilege on Sailors to Bob. The authorization graph at this interrnediate point is shown in Figure 21.2.

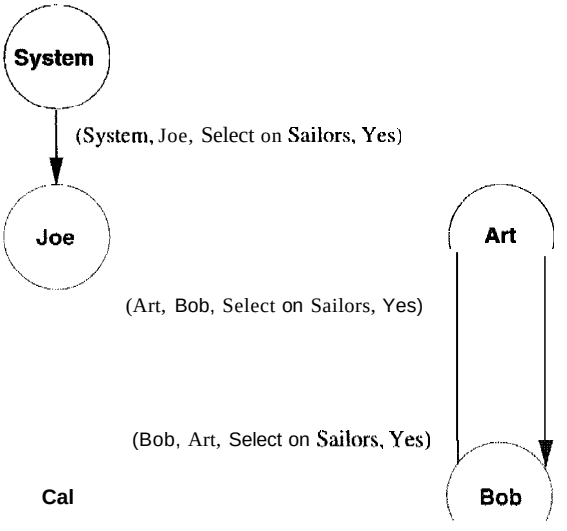


Figure 21.2 Example Authorization Graph during Revocation

rrhe graph now contains two nodes (Art and Bob) for which there are outgoing arcs with labels containing the SELECT privilege on Sailors; therefore, these users have granted this privilege. lInwever, although each node contains an incorning arc carrying the saIne privilege, *there is no such path from Systern to either of these nodes*; so these users' right to grant the privilege has been abandonecL We therefore rernove the outgoing arcs as well. In general, these nodes rnight have other arcs incident on theIn, but in this exarnplc, they now have no incident arcs. Joe is left as the only user\vith the SELECT privilege on Sailors; Art and Bob have lost their privileges.

21.3.1 Grant and Revoke on Views and Integrity Constraints

The privileges held by the creator of a view (\vitll respect to the vie\v) change over time as he or she gains or loses privileges on the underlying tables. If the creator loses a privilege held 'with the grant option, users who were given that privilege on the view lose it as \vell. There are SOIHO subtle aspects to the GRANT and REVOKE conunands when they involve views or integrity constraints. We consider senne exarnples that highlight the following irnportant points:

1. A view Inay be dropped because a SELECT privilege is revoked froIn the user who created the view.
2. If the creator of a vie"v gains additional privileges on the underlying tables, he or she autornatically gains additional privileges on the view.
3. The distinction between the REFERENCES and SELECT privileges is irnportanto

Suppose that Joe created Sailors and gave Michael the SELECT privilege on it with the grant option, and J\!Iichael then created the view YoungSailors and gave Eric the SELECT privilege on YoungSailors. Eric now defines a view called FineYoungSailors:

```
CREATE VIEW FineYoungSailors (naIne, age, rating)
        AS SELECT S.snarne, S.age, S.rating
           FROM    YoungSailors S
           WHERE   S.rating > 6
```

What happens if Joe revokes the SELECT privilege on Sailors froln l\1icha,el? Michael no longer has the authority to execute the query used to define Young-Sailors because the definition refers to Sailors. rrherefore, the view YoungSailors is dropped (i.e., destroyed). In turn, Fine'{oungSailors is dropped as well. Both view definitions are rernoved fr0111 the systcln catalogs; even if (1, rerIlorseful Joe decides to give ba,ckthe SELECT privilege on Sailors to Michael, the views are gone a11d rnust be created afresh if they are required.

On a Inore happy note, suppose that everything proceeds as just described until Eric defines FineYoungSailors; then, instead of revoking the SELECT privilege on Sailors frorll Michael, .Joe decides to also give Michael the INSERT privilege 011 Sailors. Michael's privileges on the view YoungSailors are upgraded to what he would have if he were to create the vie\v *now*. He therefore acquires the INSERT privilege on 'YourlgSailors as well. (Note that this view is updatal)le.) What about Eric? His privileges are unchanged.

Whether OI' II.ot Michael has tlle INSERT privilege 011 \roungSailors with the gra11t ()ption depends 011 whether or not .Joe gives hirn the INSERT I)rivilege OII

Sailors with the grant option. To understand this situation, consider Eric again. If Michael has the INSERT privilege on YoungSailors with the grant option, he can pass this privilege to Eric. Eric could then insert rows into the Sailors table because inserts on YoungSailors are effected by rnodifying the underlying base table, Sailors. Clearly, we do not want Michael to be able to authorize Eric to rnake such changes unless Michael has the INSERT privilege on Sailors with the grant option.

rrhe REFERENCES privilege is very different froIll the SELECT privilege, as the following exarIlple illustrates. Suppose that Joe is the creator of Boats. He can authorize another user, say, Fred, to create H,eserves with a foreign key that refers to the *bid* colurnn of Boats by giving Fred the REFERENCES privilege with respect to this colulnn. ()n the other hand, if Fred has the SELECT privilege on the *bid* colurnn of Boats but not the REFERENCES privilege, Fred *cannot* create R.eserves with a foreign key that refers to Boats. If Fred creates R,eserves with a foreign key colunlll that refers to *bid* in Boats and later loses the REFERENCES privilege on the *bid* colurnn of boats, the foreign key constraint in Reserves is dropped; however, the R,eserves table is *not* dropped.

To understand why the SQL standard chose to introduce the REFERENCES privilege rather than to siInply allow the SELECT privilege to be used in this situation, consider what happens if the definition of Reserves specified the NO ACTION option with the foreign key------.-Joe, the owner of Boats, Inay be prevented from deleting a row fronl Boats because a row in Reserves refers to this Boats row. Giving Fred, the creator of Reserves, the right to constrain updates on Boats in this rnanner goes beyond. siInply allowing hinl to read the values in Boats, which is all that the SELECT privilege authorizes.

21.4 MANDATORY ACCESS CONTROL

Discretionary access coutrollnechanislns, while generally effective, have certain \veaknesses. In particular they are susceptible to *Trojan horse* schelnes whereby a devious unauthorized user can trick an authorized user into disclosing sensitive data. For exalnple, suppose that student rrricky Dick wants to break into the grade tables of instructor Trustin Justin. IJick does the following:

- He creates a new table called MineAllMine and gives INSERT privileges on this tahle to .Justin (who is blissfully unaware of all this attention, of course).
- He rllodifies the code of SOllIe I}BlVIS application that Jllstin uses often to do a couple of additional things: first, read the Grades table, Cuid next, write the result into MineAllMine.

Then he sits back and waits for the grades to be copied into MineAllMine and later undoes the Illodifications to the application to ensure that Justin does not sOlnehow find out later that he has been cheated. Thus, despite the DBMS enforcing all discretionary access controls—only Justin's authorized code was allowed to access Grades—sensitive data is disclosed to an intruder. The fact that Dick could surreptitiously modify Justin's code is outside the scope of the DBMS's access control rnechanisrn.

NIandatory access control meehanisrns are airned at addressing such loopholes in discretionary access control. The popular rllodel for mandatory access control, called the Bell-LaPadula Illodel, is described in tenllS of objects (e.g., tables, views, rows, columns), subjects (e.g., users, prograrlls), security classes, and clearances. Each database object is assigned a *security class,* and each subject is assigned *clearance* for a security class; we denote the class of an object or subject *A* as *class(A).* The security classes in a systerll are organized according to a partial order, with a most secure class and a least secure class. For sirnplicity, we assume that there are four classes: *top secret (T8), secret (8), confidential (C), and unclassified (U).* In this system, $T8 > S > C > U$, where $A > B$ rneans that class *A* data is more sensitive than class *B* data.

The Bell-LaPadula model imposes two restrictions on all reads and writes of database objects:

1. Simple Security Property: Subject *S* is allowed to read object 0 only if *class(8)* $\geq$ *class*(()). For exarllple, a user with *TS* clearance can read a table with C clearance, but a user with C clearance is not allowed to read a table with *TS* classification.

2. *-Property: Subject *S* is allowed to write object 0 only if *class(S)* $\leq$ *class(O).* For exarllple, a user with *S* clearance can write only objects with *S* or *TS* classification.

If discretionary a,ccess controls are also specified, these rules represent additionaJ restrictions. Therefore, to read or write a database object, a user lllUst have the necessary privileges (obtained via GRANT cornrnands) *and* the security classes of the user and the object rnust satisfy the preceding restrictions. Let us consider how such a mandatory control rnech.an.isrn lui.ght h.ave foiled Tricky I)ick. rfhe Grades table could be classified as *S,* .Justin could be given clearance for *S,* and Tricky Di.ck could be given a lower clearance (*C*). Dick can create objects of only *C* or lower classification; so the table MineAllMine can have at Inos1, the classification *C.* When the application prograrIl running on behalf of ..Justin (and therefore\vith clearance *S*) tries to copy Grades into MineAllMine, it is not allowed to do so because *class(MineAllMine)* < *class(applicat'ion),* and the *-Property is violated.

21.4.1 Multilevel Relations and Polyinstantiation

rro apply Inandatory access control policies in a relational DBMS, a security class must be assigned to each database object. The objects can be at the granularity of tables, rows, or even individual colurnn values. Let us assU111e that each row is assigned a security class. This situation leads to the concept of a multilevel table, which is a table with the surprising property that users with different security clearances see a different collection of rows when they access the sarne table.

Consider the instance of the Boats table shown in Figure 21.3. Users with *S* and *TS* clearance get both rows in the answer when they ask to see all rows in Boats. A user with C clearance gets only the second row, and a user with *[J* clearance gets no rows.

bid	*bname*	*color*	Security Class
101	Salsa	Red	*S*
102	Pinto	Brown	*C*

Figure 21.3 An Instance *B1* of Boats

The Boats table is defined to have *bid* as the prirnary key. Suppose that a user with clearance C wishes to enter the row *(101, Picante, Scarlet, C)*. We have a dilemrna:

- If the insertion is perlnitted, two distinct rows in the table have key 101.
- If the insertion is not pennitted because the priInary key constraint is violated, the user trying to insert the new row, who has clearance C, can infer that there is a boat with *bid=101* whose security class is higher than C. This situation cOlnpromises the principle that users should not be able to infer any infonnation about objects that have a higher security classification.

This dilerrllla is resolved by effectively treating the security classification as part of the key. rrhus, the insertion is allo\ved to continue, and the table instance is rnodified as shown in Figure 21.4.

bid	*bna'me*	*color*	Security Class
101	'--'Salsa	Red	*S*
101	Picante	Scarlet	*C*
102	Pinto	Brown	*C*

Figure 21.4 Insta.nce 131 after Insertion

lJsers\vith clearance *C* or [1 see just the rows for Picante and Pinto, but users with clearance *S* or *TS* see all three rows. The two ro\vs with *bid*=1()1 can be interpreted in one of two ways: only the rc)\v\vith the higher classification (Salsa, with classification 8) a,ctually exists, or both exist and their presence is revealed to users according to their clearance level. The choice of interpretation is up to application developers and users.

The presence of data objects that appear to have different values to users with different clearances (for exarnple, the boat with *b'id* 101) is called polyinstantiation. If we consider security classifications associated with individual colurnns, the intuition underlying polyinstantiation can be generalized in a straightforward manner, but SOIne additional details Inust be addressed. We relnark that the rnain drawback of rnandatory access control schelnes is their rigidity; policies are set by systeIll adrninistrators, and the classification 1nechanisrns are not flexible enough. A satisfactory cornbination of discretionary and rnandatory access controls is yet to be achieved.

21.4.2 Covert Channels, DoD Security Levels

Even if a DElVIS enforces the rnandatory access control schenle just discussed, inforrnation can flow frorn a higher classification level to a lower classification level through indirect rneans, called covert channels. For exanlplc, if a transaction accesses data at rnore than one site in a distributed DBMS, the actions at the two sites 1nust be coordina,ted. The process at one site rTlay have a lower clearance (say, C) than the process at another site (say, *S*), and both processes have to agree to cOllnnit before the transaction can be conunitted. This requirernent can be exploited to pass i11fo1'rnatio11 with an *S* classification to the process with a *C* clearance: The transaction is repeatedly invoked, and the process \vith the *C* clearance always agrees to cOlIllnit, whereas the process with the *S* clearance agrees to conunit if it wants to translnit a 1 bit and does not agree if it wants to transrnit a 0 l)it.

In this (adrnittedly tortuous) lIlanllcr, infonnation with an *S* clearance can be sent to a process with a *C* clearance as a strea111 of bits. This covert cllannel is an indirect violation of the intent behind the *-Property. Additional exarnples of covert channels can be found readily in statistical databases, which we cliscuss ı'n Scetlon 21,.6.2".

DBMS vendors recently started irnplcrnenting rnandatory access control mechaniSlns (although they are not part: of the SQL standard) because the United States l)epartnlent of J)efense (1)01)) requires such support for its systems. The Dol) requirernents can be described in terrns of security levels *A*, *13*, *C*, and *D*. of \vhich *A* is the 1J10st secure and 1) is the least secure.

Current Systems: Commercial RDBMSs are available that support discretionary controls at the *C2* level and mandatory controls at the *B1* level. IBM DB2, Inforruix, Microsoft SQL Server, Oracle 8, and Sybase ASE all support SQL's features for discretionary access controL In general, they do not support lnandatory access control; Oracle offers a version of their product with support for rnandatory access control.

Level C requires support for discretionary access control. It is divided into sublevels *Cl* and *C2*; *C2* also requires sollic degree of accountability through procedures such as login verification and audit trails. Level *B* requires support for lnandatory access control. It is subdivided into levels *Bl, B2,* and *B3.* Level 132 additionally requires the identification and clirnination of covert channels. Level *B3* additionally requires 11laintenance of audit trails and the designation of a security administrator (usually, but not necessarily, the DBA). Level *A,* the most secure level, requires a n1athernatical proof that the security rnechanisrn enforces the security policy!

21.5 SECURITY FOR INTERNET APPLICATIONS

When a DBMS is accessed frorn a secure location, we can rely upon a shnple password rnechanisrn for authenticating users. IIowever, suppose our friend Sarn wants to place an order for a hook over the Internet. rrhis presents sorne unique challenges: Saln is not even a known user (unless he is a repeat custonler). Fron1 Alnazon's point of view, we have an individual asking for a book and offering to pay with a credit card registered to Saln, but is this individual really Sarn? From Sarn's point of view, he sees a fornl asking for credit card inforrnation, but is this indeed a legitirnate part of Arnazon's site, and not a rogue application designed to trick hill into revealing his credit card nurnber?

Tlhis exarnple illustrates the need for a rnore sophisticated approach to authentication than a sirnple password rnechanisrn. Encryption techniques provide the foun,dation for rnodern authentica,tion.

21.5.1 Encryption

The basic idea behind encryption is to apply an encryption algorithrn to the data, using a user-specified or IJBA-SI)Ccified encryption key. The output of the algorithrn is the encrypted version of the data. There is aJso a decryption algorithrrL ·which takes the encrypted data and a decryption key as input and then returns the original data.\Vithont the correct decryption key, the decryption algoritll111produces gibl)crish. rrhe encryption and clecryption

DES and AES: The DES standard, adopted in 1977, has a 56-bit encryption key. Over time, computers have become so fast that, in 1999, a special-purpose chip and a network of PCs were used to crack DES in under a day. The systern was testing 245 billion keys per second when the correct key was fonnd! It is estimated that a special-purpose hardware device can be built for under a 1l1iUioll dollars that can crack DES in under four hours. Despite growing concerns about its vulnerability, DES is still widely used. In 2000, a successor to DES, called the Advanced Encryption Standard (AES), was adopted as the new (syrrunetric) encryption standard. AES has three possible key sizes: 128, 192, and 256 bits. With a 128 bit key size, there are over $3 \cdot 10^{38}$ possible AES keys, which is on the order of 10^{24} Inore than the number of 56-bit DES keys. Asslllne that we could build a conlputer fast enough to crack DES in 1 second. This COIllputer would. cornpnte for about 149 trillion years to crack a 128-bit AES key. (Experts think the universe is less than 20 billion years old.)

algorithrns thernselves are assunled to be publicly known, but one or both keys are secret (depending upon the encryption scheme).

In symmetric encryption, the encryption key is also used as the decryption key. The ANSI Data Encryption Standard (DES), which has been in use since 1977, is a well-known exarnple of syllunetric encryption. It uses an encryption algorithrn that consists of character substitutions and pernlutations. The nlain weakness of synunetric encryption is that all authorized users rnust be told the key, increasing the likelihood of its becorning known to an intruder (e.g., by sirnple Inllnan error).

Another approach to encryption, called public-key encryption, has becorne increasingly popular in recent years. The encryption scheniC proposed by Hjvest, Sharnir, and Adlernan, called RSA, is a well-known exarnple of public-key encryption. Each authorized user has a public encryption key, known to everyone, and a private decryption key, known only to hini or her. Since the private decryption keys are known only to their owners, the weakness of 1)ES is avoided.

A central issue for public-key encryption is how encryption and decryption keys are chosen. Technically, public-key encryption algorithrns rely on the existence of one-way functions, whose inverses are cornplltationally very hard to deterrnine. rrhe RSA algoritllIn, for example, is based on the observation that, although checking whether a given nurnber is prirne is easy, deterrnining the prirne factors of a nonprime nurnber is extrernely hard. (I)eterlnining the

Why RSA Works: The essential point of the scherne is that it is easy to compute d given e, p, and q, but *very* hard to cornpute d given just e and L. In turn, this difficulty depends on the fact that it is hard to determine the prime factors of L, which happen to be p and q. *A caveat:* Factoring is widely believed to be hard, but there is no proof that this is so. Nor is there a proof that factoring is the only way to crack RSA; that is, to comp d frolll e and L.

prirne factors of a nurnber with over 100 digits can take years of CPlJ tirne on the fastest available COIllputers today.)

We now sketch the idea behind the RSA algorithrn, assuming that the data to be encrypted is an integer I. To choose an encryption key and a decryption key for a given user, we first choose a very large integer L, larger than the largest integer we will ever need to encode.[1] We then select a nUl1lber e as the encryption key and cornpute the decryption key d based on e and L; how this is done is central to the approach, as we see shortly. Both L and e are l1lade public and used by the encryption algorithrn. However, d is kept secret and is necessary for decryption.

- The encryption function is S $Ie \bmod L$.
- The decryption function is 1 $Sd \bmod L$.

We choose L to be the product of two large (e.g., 1024-bit), distinct prirne nurnbers, $p * q$. The encryption key e is a randornly chosen nlunber between 1 and L that is relatively prirne to $(p - 1) * (q - 1)$. The decryption key d is cornputed such that $d * e = 1 \bmod ((p-1)*(q-1))$. Given these choices, results in nurnber theory can be used to prove that the decryption function recovers the original ruessage frorll its encrypted version.

A very irnportant property of the encryption and decryption algoritluns is that the roles of the encryption and decryption keys can be reversed:

$$decrypt(\mathrm{d}, (encrypt(\mathrm{e}, \mathrm{I}))) = \mathrm{I} = decrypt(\mathrm{c}, (encrypt(\mathrm{d}, \mathrm{I})))$$

Since In.any protocols rely on this property, we henceforth sirnply refer to public and private keys (since both keys can be used for encryption as well as decryption).

[1] A message that is to be encrypted is decomposed into blocks such that each block can be treated as an integer less tha.n L.

While we introduced encryption in the context of authenticatioll, we note that it is a fundaIllental tool for enforcing security. A DBMS can use *encryption* to protect inforrnation in situations where the norrnal security rnechanisrns of the DBlVIS are not adequate. For exarnple, an intruder rnay steal tapes containing SOUIC data or tap a conunu.nieation line. By storing and transrnitting data in an encrypted forIn, the DBMS ensures that such stolen data is not intelligible to the intruder.

21.5.2 Certifying Servers: The SSL Protocol

Suppose we associate a public key and a decryption key with Alnazon. Anyone, say, user Sam, can send Alnazon an order by encrypting the order using Arnazon's public key. ()nly Arnazon can decrypt this secret order because the decryption algorithrn requires Arnazon's private key, known only to Arnazon.

This hinges on 8arn's ability to reliably find out Arnazon's public key. A number of cornpanies serve as certification authorities, e.g., Verisign. Arnazon generates a public encryption key e_A (and a private decryption key) and sends the public key to Verisign. Verisign then issues a certificate to Arnazon that contains the following inforrnation:

$$\langle \textit{Verisign Arnazoin, htl;P8:/www.arnazon.corn,} \; e_A \rangle$$

The certificate is encrypted using Verisign's own *private* key, which is known to (i.e., stored in) Internet Explorer, Netscape Navigator, and other browsers.

When 8an1 carnes to the Amazon site and wants to place an order, his browser, running the SSL protocol,[2] asks the server for the Verisign certificate. The browser then validates the certificate by decrypting it (using ·Verisign's public key) and checking that the result is a certificate with the HaIne Verisign, and that theURL it contains is that of the server it is talking to. (Note that an atternpt to forge a certificate will fail because certificates are encrypted using Verisign's private key, whieh is known only to Verisign.) Next, the browser generates a random session key, encrypt it using Arnazon's public key (which it obtained frorn the validated certificate anel therefore trusts), and sends it to the Amazon server.

Frorn this point on, the Arnazon server and the browser can use th.c session key (which both know and are confident tliat only they know) and a *symmetric* encrypticHl protc)collike AES or IJES to exchange securely encrypted rnessages: Messages are encrypted by the sender anel decrypted by the receiver using the sa,HIe session key. rrhe encrypted Inessages travel over the Internet and rnay be

[2] A browser uses the SSL protocol if the target URL begins with *https*.

intercepted, but they cannot be decrypted without the session key. It is useful to consider why \ve need a session key; after all, the bro\vser could sirnply have encrypted 8a1n's original request using Arnazon's public key and sent it securely to the Arnazon server. The reason is that, without the session key, the Amazon server has no way to securely send infonnation back to the bro\vser. A further advantage of session keys is that syrnrnetric encryption is cOlnputationally nluch faster than public key encryption. The session key is discarded at the end of the session.

Thus, 8aIn can be assured that only Alnazon can see the inforrnation he types into the fonn shown to hirn by the AIuazon server and the inforrnation sent back to hiln in responses froIn the server. However, at this point, Amazon has no assurance that the user running the browser is actually Sanl, and not SOlneone who has stolen Sarn's credit card. l-'ypically, rnerchants accept this situation, which also arises when a custoIner places an order over the phone.

If we want to be sure of the user's identity, this can be accoIuplished by additionally requiring the user to login. In our exarnple, 8arn 11lUSt first establish an account with Alnazon and select a password. (Sam's identity is originally established by calling hiln back on the phone to verify the account inforrnation or by sending elnail to an elnail address; in the latter case, all we establish is that the owner of the account is the individual with the given clnail address.) Whenever he visits the site and Anlazon needs to verify his identity, AInazon redirects hinl to a login fo1'1n *after* using SSL to establish a session key. The password typed in is transrnitted securely by encrypting it with the session key.

()ne rcrnaining drawback in this approach is that Arnazon now kno\vs Sarn's credit card nlunber, and he rnust trust Alnazon not to rnisuse it. The Secure Electronic Transaction protocol addresses this lirnitation. Every custolner rnust now obtain a certificate, with his or her own private and public keys, and every transaction involves the Alnazon server, the cust(nner's browser, and the server of a trusted third party, such as Visa for credit card transactions. The basic idea is that the browser encodes non-credit caTd inforrnation using AInazon's public key and the credit ca.rd infonnation using Visa's public key and sends these to the AJnazon server, which for"vards the credit card inforrnation (which it cannot decrypt) to the Visa server. If the Visa server a,pproves the inforrnation, the transa,ction goes through.

21.5.3 Digital Signatures

Suppose tllat ,Elnlcr, who works for ArnazoIl, a,nd Betsy, who works for McGrawlIill,need to COllllll1Unicate with each other about inventory. Public key encryption can be used t() create digital signatures for rnessages. rrhat is, rnessages

can be encoded in such a way that, if Elmer gets a Inessage supposedly fr(nTI Betsy, he can verify that it is fronl Betsy (in addition to being able to decrypt the rnessage) and, further, *prove* that it is froIn Betsy at McGraw-lIill, even if the ll1cssage is sent froIn a IIotrnail account when Betsy is traveling. Sirnilarly, Betsy can authenticate the originator of Inessages froln Ellner.

If Ellner encrypts Inessages for Betsy using her public key, and vice-versa, they can exchange inforrnation securely but cannot authenticate the sender. Sorueone who wishes to irnpersonate Betsy could use her public key to send a rnessage to EIrner, pretending to be Betsy.

A clever use of the encryption sche111e, however, allows Elmer to verify whether the rnessage was indeed sent by Betsy. Betsy encrypts the rnessage using her *private* key and then encrypts the result using Elrner's public key. When Ellner receives such a 111essage, he first decrypts it using his private key and then decrypts the result using Betsy's public key. rrhis step yields the original un-encrypted rnessage. Furthermore, Ehner can be certain that the rnessage was composed and encrypted by Betsy because a forger could not have known her private key, and without it the final result would have been nonsensical, rather than a legible 111essage. Further, because even Elmer does not know Betsy's private key, Betsy cannot clairn that Ehner forged the ruessage.

If authenticating the sender is the objective and hiding the rnessage is not im-portant, we can reduce the cost of encryption by using a message signature. A signature is obtained by applying a one-way function (e.g., a hashing schelne) to the rnessagc and is considerably sInaHer. We encode the signature as in the basic digital signature approach, and send the encoded signature together with the full, unencoded 111cssage. rrhe recipient can verify the sender of the signa-ture as just described, and validate the I11essage itself by applying the one-way function and cOlnparing the result with the signature.

21.6 ADDITIONAL ISSUES RELATED TO SECURITY

Security is a l)road topic, and our coverage is necessarily lirnited. 'rhis section briefly touches on sorne additional irnportant issues.

21.6.1 Role of the Database Administrator

rrhe database administrator (IJBA) plays an irnportant role in enforcing the security-related aspects of a database design. In conjunction with the o\vners of the data, the I)BA aJso COlltributes to developing a security policy. The I)BA has a special i:1,ccount, which we call the systenl account, and is responsible

for the overall security of the systeru. In particular, the DBA. deals with the follo\ving:

1. Creating New Accounts: Each new user or group of users Blust be assigned an authorization ID and a password. Note that application prograIns that access the database have the saIne authorization ID as the user executing the prograill.
2. Mandatory Control Issues: If the DBMS supports rnandatory control·-···-S0111e custornized systeIns for applications with very high security requirernents (for exarllple, rnilitary data) provide such support—the DBA lllUst assign security classes to each database object and assign security clearances to each authorization ID in accordance with the chosen security polICY·

The DBA is also responsible for rnaintaining the audit trail, which is essentially the log of updates with the authorization ID (of the user executing the transaction) added to each log entry. This log is just a Ininor extension of the log mechanislll used to recover from crashes. Additionally, the DBA may choose to rnaintain a log of *all* actions, including reads, perfornled by a user. Analyzing such histories of how the DBMS was accessed can help prevent security violations by identifying suspicious patterns before an intruder finally succeeds in breaking in, or it can help track down an intruder after a violation has been detected.

21.6.2 Security in Statistical Databases

A statistical database contains specific inforrnation on individuals or events but is intended to perlnit only statistical queries. For exarnple, if we rnailltained a statistical database of inforrna,tion about sailors, we would allow statistical queries about average ratings, rnaxirnurn age, and so on, but not queries about individua.l sailors. Security in such databases poses new probleurs because it is possible to infer protected inforrnation (such as a sailor's rating) frorn answers to perrnitted statistical queries. Such inference opportunities represent covert channels that can cornprornise the security policy of the database.

Suppose that sailor Sneaky Pete wants to kncrw the rating of .A.clrniral Ho1'ntooter, the esteemed chairrnarl of the sailing clu1), and happens to kno\v that IIorntooter is the oldest sailor in the club. Pete repeatedly asks queries of the forln "How InClny sailors are there whose age is greater than X?" for various values of X, until the answer is 1. Obviously, this sa,iIor is lIorntooter, the oldest sailor. Note that each of these queries is a valicl statistical query and is permitted. Let the value of X at this point be, say, 65. Pete no\v asks the query, "What is the nraxirnurn rating of all sailors \vhose age is greater than

65?" Again, this query is pennitted because it is a statistical query. However, the answer to this query reveals J101'ntooter's rating to Pete, and the security policy of the database is violated.

One approach to preventing such violations is to require that each query rnust involve at least S01ne lnininuull nUluber, say, N, of l'()\VS. With a reasonable choice of N, Pete \vould not be able to isolate the inforrnation about 1101'ntooter, because the query about the maximum rating would fail. rrhis restriction, however, is easy to overCOIne. By repeatedly asking queries of the forIII, "How ruany sailors are there whose age is greater than X?" until the systenl rejects one such query, Pete identifies a set *ofN* sailors, including Florntooter. Let the value of X at this point be 55. Now, Pete can ask two queries:

- "What is the SlUll of the ratings of all sailors whose age is greater than 557" Since N sailors have age greater than 55, this query is perrnitted.
- "What is the SUIll of the ratings of all sailors, other than l1orntooter, whose age is greater than 55, and sailor Pete?" Since the set of sailors whose ratings are added up now includes Pete instead of Horntooter, but is otherwise the sallle, the rnunber of sailors involved is still N, and this query is also pennitted.

From the answers to these two queries, say, AI and A_2, Pete, who knows his rating, can easily calculate Horntooter's rating as $AI - A_2$ + *.Pete'8 rating.*

Pete succeeded because he was able to ask two queries that involved Illany of the sarne sailors. 'The nurnber of rows exalnined in corllrnon by two queries is called their intersection. If a limit were to be pla,ced on the alIlount of intersection perrnitted bet\veen any two queries issued by the same user, Pete could be foiled. Actually, a truly fiendish (and patient) user can generally find out inforruation about specific individuals even if the systcrn places a, rniniruurn nUlnber of ro\vs bound *(N)* and a rnaxirnurn intersection bound (M) on queries, hut the n1.11n])er of queries required to do this gro\vs in proportion to N/M. We can try to additionally lirnit the total nUlnbe1' of queries that a user is allowed to ask. but two users could still conspire to breach security. By Il1aintaining a log of all activity (including read-only accesses), such query patterns can be detected, icleally before a security violation occurs. This discussion should rnake it clear, however, that security in statistical databases is difficult to enforce.

21.7 DESIGN CASE STUDY: THE INTERNET STORE

We return to our case study and our friends at DBI)udes to consider security issues. 'There are three groups of users: custolners, employees, and the owner of the l>ook shop. (()f course, there is also the database adrninistrator, who

has universal access to all data and is responsible for regular operation of tlle database systcrTl.)

The owner of the store has full privileges on all tables. Custorners can query the Books table and place orders online, but they should not have access to other customers' records nor to other c11storne1'8' orders.DBDudes restricts access in two ·ways. First, it designs a simple Web page with several fonus similar to the page shown in Figure 7.1 in Chapter 7. This allo\vs custo1ners to subrnit a s111all collection of valid requests without giving tho1n the ability to directly access the underlying DBMS through an SQL interface. Second, I)B.Dudes uses the security features of the DBMS to li111it access to sensitive data.

rIhe \vebpage allows custorners to query the Books relation by ISBN nU111bcr, narne of the author, and title of a book. The webpage also has two buttons. The first button retrieves a list of all of the custolIler's orders that are not completely fulfilled yet. l'he second button displays a list of all cornpleted orders for that custorner. Note that custolllers cannot specify actual SQL queries through the Web but only fill in SCHne pararneters in a forn1 to instantiate an autonlatically generated SQL query. All queries generated through fonll input have a WHERE clause that includes the *cid* attribute value of the current custo1ner, and evaluation of the queries generated by the two buttons requires knowledge of the custolller identification nUlnber. Since all users have to log on to the website before browsing the catalog, the business logic (discussed in Section 7.7) lllust 1naintain state inforrnation about a custoDler (i.e., the Cllstorner identification nUlnber) during the custorner's visit to the website.

The second step is to configure the database to lirnit access according to each user group's need to know. DBI)udes creates a special customer account that has the following privileges:

```
SELECT ON Books, NewOrders, ()ldOrders, NewOrderlists, OldOrderlists
INSERT ON New()rders, OldOrders, New()rderlists, ()ld()rderlists
```

Ernployees should be able to acid new books to the catalog, upda,te the quantity of a book in stock, revise custorner orders if necessary, and update all custorner inforrnation *except the credit card information.* In fact, ernployees should not even be able to see a custorner's credit card nurnber. 1]lcreforc,DBDucles creates the following view:

```
CREATE VIEW CustomerInfo (cid,cnarnc,address)
        AS SELECT C.cid, C.cname, C.(l.cldress
           FROMCllstolners C
```

I)BI)udes gives the employee account the follc)\ving privileges:

```
SELECT ON CustomerInfo, Books,
          NewOrders, ()IdOrders, NewOrderlists, Old()rderlists
INSERT ON Cllstornerlnfo, Books,
          Nc\vOrders, 01dC)rders, NewOrderlists, ()ldOrderlists
UPDATE ON CustolnerInfo, Books,
          New()rders, OldOrders, NewOrderlists, Old()rderlists
DELETE ON Books, NewOrders, OldOrders, NewOrderlists, ()ldOrderlists
```

Observe that ernployees can rnodify Custornerlnfo and even insert tuples into it. This is possible because they have the necessary privileges, and further, the view is updatable and insertable-into. While it seerns reasonable that elllployees can update a custorner's address, it does sceln odd that they can insert a t1lple into Cllstornerlnfo even though they cannot see related infonna.tion about the custorner (i.e., credit card nurnber) in the Cllstorners tahle. The reason for this is that the store wants to be able to take orders 1'1'0111 first-tirne custorners over the phone without asking for credit card inforrnation over the phone. Ernployees can insert into Custornerlnfo, effectively creating a new Custoillers record without credit card inforluation, and custorners can subsequently provide the credit card nurnber through a Web interface. (Obviously, the order is not shipped until they do this.)

In addition, there are security issues when the user first logs on to the website using the cllstolner identification nUlnber. Sending the nUlnber unencrypted over the Internet is a security hazard, and a secure protocol such as SSL should be used.

Cornpanies such as CyberCash and DigiCash offer electronic conunerce pay-rIlcnt solutions, even inclu.ding *electronic cash.* Discussion of how to incorporate such techniques into the website are outside the scope of this book.

21.8 REVIEW QUESTIONS

Answers to the review questions can be founel in the listed sections.

- What are tIle 1n,ain objectives in designing a secure datal)ase application? Explain the te1'rns *secrecy, integrity, availability,* and *authentication.* (Section 21.1)
- Explain the terms *security policy* and *security mechanism* arid how tllCy are related. (Section 21.1)
- What is the Blain idea behind *discretionary access control?* What is the idea behind *mandatory access control?* What are the relative merits of these two approaches? (Section 21.2)

- Describe the privileges recognized in SQL? In particular, describe SELECT, INSERT, UPDATE, DELETE, and REFERENCES. For each privilege, indicate \vho acquires it automatically on a given table. (Section **21.3)**
- I-Io\v are the owners of privileges identified? In particular, discuss *authorization ID8* and *roles.* (Section 21.3)
- What is an *authorization graph?* Explain SQL's GRANT and REVOKE eOl11-mands in terrns of their effect on this graph. In particular, discuss what happens when users pass on privileges that they receive frorn sorneone else. (Section **21.3)**
- Discuss the difference between having a privilege on a table and on a vie\v defined over the table. In particular, how can a user have a privilege (say, SELECT) over a view 'without also having it on all underlying tables? Who 11111st have appropriate privileges on all underlying tables of the view? (Section 21.3.1)
- What are *objects, subjects, security classes,* and *cleaTances* in rnandatory access control? IJiscuss the Bell-LaPadula restrictions in tenns of these concepts. Specifically, define the *simple security property* and the **-pToperty.* (Section 21.4)
- What is a *Trojan horse* attack and how can it cOlnprornise discretionary access control? Explain how lnandatory a,ccess control protects against Trojan horse attacks. (Section 21.4)
- What do the tenns *multilevel table* and *polyinstantiation* mean? Explain their rela.tionship, a,nd how they arise in the context of lnandatory access control. (Section 21.4.1)
- What are *covert channels* and how can they arise when both discretionary and luandatory access controls are in place? (Section **21.4.2)**
- Discuss the I)oD security levels for database systclns. (Section 21.4.2)
- Explain why a sirnple password rnechanisrn is insufficient for authentication of users who access a database renJotely, say, over the Internet. (Section 21.5)
- What is the difference between *symmetric* and *public-key encryption?* Give examples of well-known encryption algoritluns of both ki.lldso What is the rnain weakness of synunetric encryption and how is this addressed in public-key encryption? (Section 21.5.1)
- 1)i8c118s the choice of encrYIltion and decryption keys in public-key en.cryption and how they are llsed to encrypt and decrypt data. Explain the role of *one-way functions.* What H.ssurance do\ve h.ave that the RSA scheme cannot be cornprornised? (Section 21.5.1)

- What are *certification authorities* and why are they needed? Explain how *certificates* are issued to sites and validated by a bro\vser using the *SSL protocol*; discuss the role of the *session key.* (Section 21.5.2)
- If a user connects to a site using the SSL protocol, explain \vhy there is still a need to login the uscr. Explain the use of SSL to protect pass\vords and other sensitive infonnation being exchal1ged. What is the *secure electronic transaction protocol*? What is the added value over SSL? (Section 21.5.2)
- A *digital signature* facilitates secure exchange of rnessages. Explain what it is and how it goes beyond sirnply encrypting rnessages. Discuss the use of *message signatures* to reduce the cost of encryption. (Section 21.5.3)
- What is the role of the database achninistrator with respect to security? (Section 21.6.1)
- Discuss the additional security loopholes introduced in *statistical databases.* (Section 21.6.2)

EXERCISES

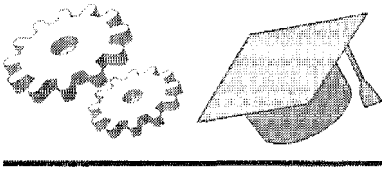
Exercise 21.1 Briefly answer the following questions:

1. Explain the intuition behind the two rules in the Bell-LaPadulamodel for rnandatory access control.
2. Give an exarnple of how covert channels can be used to defeat the Bell-LaPadula rnodel.
3. Give an exarnple of polyinstantiation.
4. Describe a scenario in whichrnandatory access controls prevent a breach of security that cannot be prevented through discretionary controls.
5. Describe a scenario in which discretionary access controls are required to enforce a seCll-rity policy that cannot be enforced using only mandatory controls.
6. If a DBMS already supports discretionary and Jnandatory access controls, is there a need for encryption?
7. Explain the need for each of the following lirnits in a statistical database systern:
   (a) A maxirnurn on the munber of queries a user can pose.
   (b) A rninirnUln on the munber of tuples involved in ans\vering a query.
   (c) A maximurn on the intersection of two queries (i.e., on the number of tuples that both queries exarnine).
8. Explain the use of an audit trail, with special reference to a statistical database system.
9. \,Vllat is the role of the DBA with respect to security?
10. Describe AES and its relationship to DES.
11. What is public-key encryption? How does it differ frorn the encryption approach taken in the Data Encryption Standard (DES), and in what ways is it better than DES?

12. Explain how a company offering services on the Internet could use encryption-based techniques to lllake its order-entry process secure. Discuss the role of DES, A.ES, SSL, SET, and digital signatures. Search the Web to find out IllOre about related techniques such as *electronic cash.*

Exercise **21.2** You are the DBA for the VeryFine Toy Cornpany and create a relation called Employees with fields *enam,e, dept,* and *salary.* For authorization reasons, you also define views EmployeeNallIes (with *ena:rne* as the only attribute) and DeptInfo with fields *dept* and *avgsalary.* The latter lists the average salary for each departrnent.

1. Show the view definition statements for EnlployeeNames and Deptlnfo.
2. What privileges should be granted to a user who needs to know only average departn1ent salaries for the Toy and CS departments?
3. You want to authorize your secretary to fire people (you will probably tell hill1 whorn to fire, but you want to be able to delegate this task), to check on who is an elllployee, and to check on average department salaries. What privileges should you grant?
4. Continuing with the preceding scenario, you do not want your secretary to be able to look at the salaries of individuals. Does your answer to the previous question ensure this? Be specific: Can your secretary possibly find out salaries of *some* individuals (depending on the actual set of tuples), or can your secretary always find out the salary of any individual he wants to?
5. You want to give your secretary the authority to allow other people to read the EUlployeeNames view. Show the appropriate conI111and.
6. Your secretary defines two new views using the EnIployeeNarnes view. The first is called AtoRNames and simply selects names that begin with a letter in the range A to R. The second is called HowManyNan1es and counts the number of narnes. You are so pleased with this achievement that you decide to give your secretary the right to insert tuples into the EnlployeeNan1es view. Show the appropriate cOllunand and describe 'what privileges your secretary has after this cornrnand is executed.
7. Your secretary allows Todd to read the Erl1ployeeNarnes relation and later quits. You then revoke the secretary's privileges. \Vhat happens to Todd's privileges?
8. Give an exarnple of a view update on the preceding schelna that cannot be illlplernentecl through updates to Erl1ployees.
9. You decide to go on an extended vacation, and to rnake sure that ernergencies can be handled, you want to authorize your boss Joe to read and modify the Employees relation and the ErIlployeeNalnes relation (and Joe lllust be able to delegate authority, of course, since he is too far up the managernent hierarchy to actually do any \vork). Show the appropriate SQL staternents. Can Joe read the Deptlnfo view?
10. After returning frorn your (wonderful) vacation, you see a note from Joe, indicating that he authorized his secretary Mike to read the Ernployees relation. You \vant to revoke Mike's SELECT privilege on Ernployees, but you do not \vant to revoke the rights you gave to Joe, even teruporarily. Can you do this in SQL?
11. Later you realize that Joe has been quite busy. He has defined a view called AllNarnes using the view ErnployeeNames, defined another relation called StaffNarnes that he has access to (but you cannot access), and given his secretary Mike the right to read from the AllNames view. Mike has passed this right on to his friend Susan. You decide that, even at the cost of annoying Joe 1)y revoking Borne of his privileges, you sirnply have to take away Mike (\nd Susarl's rights to see your data. What REVOKE staternent \vould you execute? What rights does Joe have on Ernployees after this statement is executed? What views are dropped as a consequence?



22

PARALLEL AND DISTRIBUTED DATABASES

- What is the rnotivation for parallel and distributed DBMSs?
- What are the alternative architectures for parallel database systellls?
- How are pipelining and data partitioning used to gain parallelism?
- How are dataflow concepts used to parallelize existingsequential code?
- What are alternative architectures for distributed DBMSs?
- How is data distributed across sites?
- How can we evaluate and optimize queries over distributed data?
- What are the nlerits of synchronous vs. asynchronous replication?
- How are transactions Inanaged in a distributed environment?
- Key concepts: parallel DBMS architectures; perfonnance, speed-up and scale-up; pipelined versus data-partitioned parallelism, block-ing; partitioning strategies; dataflow operators; distributed DBMS architectures; heterogeneous systernsj gateway protocols; data distri-bution, distributed catalogs; sernijoins, data shipping; synchronous versus asynchronous replication; distributed transactions, lock nlan-agcrnent, deadlock detection, two-phase ccnnlnit, Presurned Abort

No rnan IS an island, entire of itself; every Tnan IS a pIece of the contirlcnt, a part of the rnain.

—JohnDonne

In this chapter we look at the issues of parallelism and data distribution in a DBMS. We begin by introducing parallel and distributed database systcrIls in Section 22.1. In Section 22.2, we discuss alternative hardwa,re configurations for a parallel DBMS. In Section 22.3, we introduce the concept of data partitioning and consider its influence on parallel query evaluation. In Section 22.4, we show how data partitioning can be used to parallelize several relational operations. In Section 22.5, we conclude our treatrnent of parallel query processing with a discussion of parallel query optirnization.

'The rest of the chapter is devoted to distributed databases. We present an overview of distributed databases in Section 22.6. We discuss sorne alternative architectures for a distributed DBMS in Section 22.7 and describe options for distributing data in Section 22.8. We describe distributed catalog rnanagernent in Section 22.9, then in Section 22.10, we discuss query optirnization and evaluation for distributed databases. In Section 22.11, we discuss updating distributed data, and fina.lly, in Sections 22.12 to 22.14 we describe distributed transaction ruanagernent.

22.1 IN"TRODUCTION

We have thus far considered centralized database rnanageruent systerns in which all the data is luaintained at a single site and assumed that the processing of individual transactions is essentially sequential. One of the most irnportant trends in data.bases is the increased use of parallel evaluation techniques and data, distribution.

A parallel database system seeks to irnprove perforruance through parallelization of various operations, such as loading data, building indexes, and evaluating queries. Although data may be stored in a distributed fashion in such a systcrn, the distribution is governed. solely by perfon.nance considerations.

In a distributed database systenl, data, is physically stored across several sites, and each site is typically rnanaged by a DBMS capable of running il1-dependent of the 01:11e1' sites. rrhe location of data itenlS and the degree of autonorny of iJldividual sites have a significant irnpa,ct on all aspects of the system, including query optirnization and processing, concurrency control, and recovery. In contrast to parallel databases, the distribution of data is governed by factors such as locaJ ownership and increased a,vailability, in addition to perforlnance issues.

While parallelism is 1110tivated l)y performance consideratiolls, several distinct issues rnotivate data distribution:

- Increased AvailabHity: If a site containing a relation goes down, the relation continues to be available if a copy is Inaintained at another site.
- Distributed Access to Data: An organization Inay have branches in several cities. Although analysts may need to access data corresponding to different sites, we usually find locality in the access pa,tterns (e.g., a bank lnanager is likely to look up the accounts of custorners at the local branch), and this locality can be exploited by distributing the data accordingly.
- Analysis of Distributed Data: Organizations \vant to examine all the data available to thern, even when it is stored across rnultiple sites and on lllultiple database systerns. Support for such integrated access involves nlany issues; even enabling access to widely distributed data can be a challenge.

22.2 ARCHITEC"rURES FOR PARALLEL DATABASES

The basic idea behind parallel databases is to carry out evaluation steps in parallel whenever possible, and there are rnany such opportunities in a relational DBMS; databases represent one of the lnost successful instances of parallel cornputing.

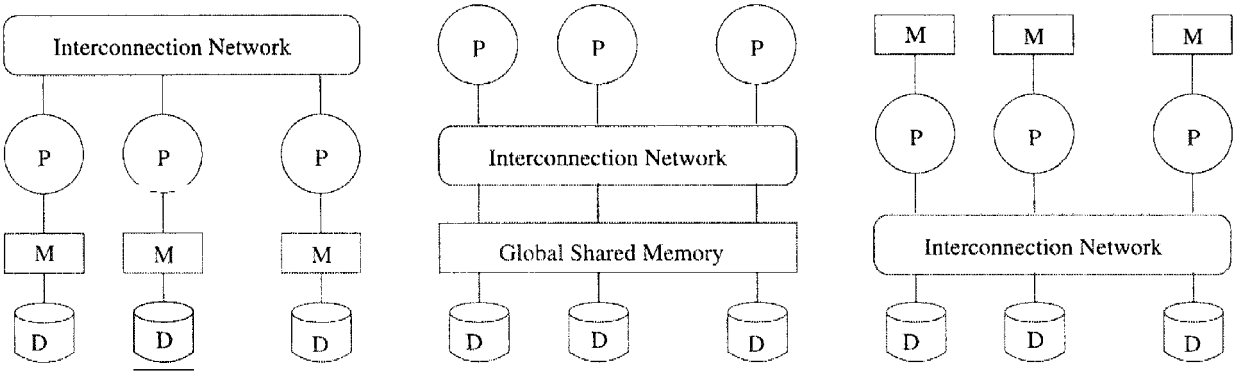


Figure 22.1 PhysicaJ Architectures for Parallel Da.tabase Systems

Three luain architectures have been proposed for building parallel DBIVISs. In a shared-IuerIlory SystCIll, Inultiple CPU·s are attached to an interconnection net\vork and can access a cornrl1on region of rnain lnelilory. In a shared-disk s:ysten.1, each CPU has a private rnelnory and direct access to all disks through an interconnection network. In a shared-nothing system, each CPTJ has local rnain lnelnory and disk space, but no two CP1Js can access the sarne storage area; all cOHununication between CP1Js is tllrough a 11etwork connection.. rrhe three architectures are illustrated in Figure 22.1.

The shared-rnernory architecture is closer to a conventional machine, and Illany conunercial database systerns have been ported to shared lnernory platfornlS \vith relative ease. Communication overhead is low, because lnain rncIIlory can be used for this purpose, and operating systern services can be leveraged to utilize the additional CPUs. Although this approach is attractive for achieving rnoderate parallelism—a few tens of CPlJs can be exploited in this fashion- Inernory contention bec01nes a bottleneck as the nurnber of CPUs increases. rfhe shared-disk architecture faces a sirnilar problcrn because large a1110nnts of data are shipped through the interconnection network.

The basic problern with the shared-111Crrlory and shared-disk architectures is interference: As Inore CPUs are added, existing CPUs are slowed down because of the increased contention for mClllory accesses and network bandwidth. **It** has been noted that even an average 1 percent slowdown per additional CPU 1nea11S that the rnaxirnum speed-up is a factor of 37, and adding additional CPlJs actually slows down the systern; a systenl with 1000 CPUs is only 4 percent as effective as a *single-CPU* systern! This observation has rllotivated the develop-rnent of the shared-nothing architecture, which is now widely considered to be the best architecture for large parallel database systems.

rrhe shared-nothing architecture requires rnore extensive reorganization of the DBNIS code, but it has been shown to provide linear speed-up, in that the tilne taken for operations decreases in proportion to the increase in the nUlnber of CPlJs and disks, and linear scale-up, in that perforrnance is sustained if the nurnber of CPUs and disks are increased in proportion to the arnount of data. Consequently, ever-rnore-powerful parallel database systcrns can be built by taking advantage of rapidly irnproving perforrllance for single-CPU systelns and connecting as rnany CPUs as desired.

Speed-up and scale-up are illustrated in Figure 22.2. 'The speed-up curves show how, for a fixed database size, Inore transactions can be executed l)cr second by adding CPUs. The scale-up curves show how adding Inorc resources (in the forln of CPlJs) enables us to process larger problerns. rrhe first scale-up graph Incasures the nurnber of transactions executed per second as the clatabase size is increased and the nurnber of CPlJs is correspondingly increased. Arl alternative way to Ineasure scale-up is to consider the time taken per transaction as r1101'e CPUs are added to process an increasing nurnber of transactions per second; the goal here is to sustain the response tirne per transaction.

22.3 PARALLEL QUERY EVALUATION'

In this section, we discuss parallel evaluation of a relational query in a DBMS with a shared-nothing architecture. While it is possible to consicler parallel

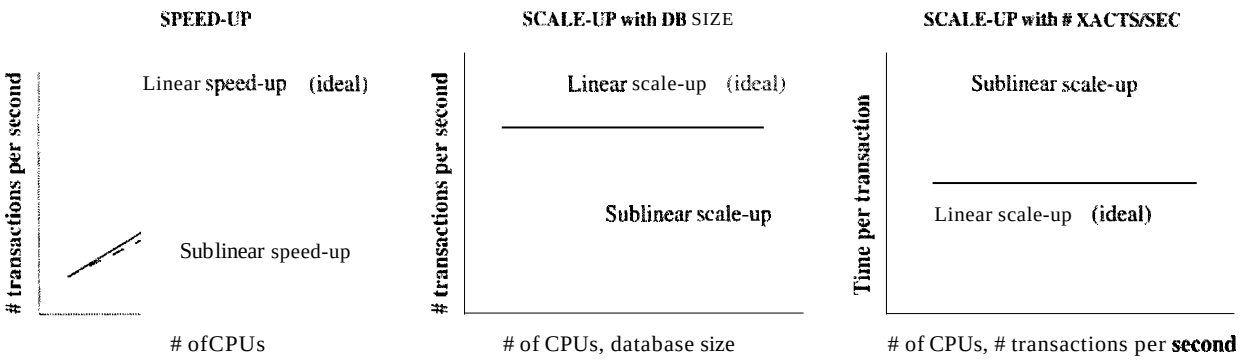


Figure 22.2 Speed-up and Scale-up

execution of rnultiple queries, it is hard to identify in advance which queries will run concurrently. So the ernphasis has been on parallel execution of a single query.

A relational query execution plan is a graph of relational algebra operators, and the operators in a graph can be executed in parallel. If one operator consurnes the output of a second operator, we have pipelined parallelism (the output of the second operator is worked on by the first operator as soon as it is generated); if not, the two operators can proceed esseptially independently. An operator is said to block if it produces no output until it has conSUllled all its inputs. Pipelined parallelisrn is lirnited by the presence of operators (e.g., sorting or aggregation) that block.

In addition to evaluating different operators in parallel, we can evaluate each individual operator in a query plan in a parallel fashion. rrhe key to evaluating an operator in pa,rallel is to *partition* the input data; \ve can then work on each partition in parallel and cornbine the results. This approach is called data-partitioned parallel evaluation. By exercising sorne care, existing code for sequentially evaluating relational operators can be ported easily for data-partitioned parallel evaluation.

An inlportant observation, which explains why shared-nothing parallel database systelns have been very successful, is that database query evaluation is very amenable to data-partitioned parallel evaluation. The goal is to nlinirnize data shipping by paTtitioning the data and structuring the algoritluns to do lnost of the processing at individual processors. (We lIse *processor* to refer to a CPU together with its local disk.)

We now consider data paxtitioning and parallelization of existing operator evaluation code in rnore detail.

22.3.1 Data Partitioning

Partitioning a large dataset horizontally across several disks enables us to exploit the I/O banchvidth of the disks by reading and writing theln in parallel. rrhere are several ways to horizontally partition a relation.vVe can assign tuples to processors in a round-robin fashion, we can use hashing, or we can assign tuples to processors by ranges of field values. If there are *n* processors, the 'ith tuple is assigned to processor *i rnodn* in round-robin partitioning. Recall that round-robin partitioning is used in RAID storage systelTIS (see Section 9.2). In hash partitioning, a hash function is applied to (selected fields of) a tuple to deternline its processor. In range partitioning, tuples are sorted (conceptually), and *n* ranges are chosen for the sort key values so that each range contains roughly the SalTIe nurnber of tuples; tuples in range i are assigned to processor i.

Round-robin partitioning is suitable for efficiently evaluating queries that access the entire relation. If only a subset of the tuples (e.g., those that satisfy the selection condition *age* = 20) is required, hash partitioning and range partitioning are better than round-robin partitioning because they enable us to access only those disks that contain rnatching tuples. (Of course, this statement assumes that the tuples are partitioned on the attributes in the selection condition; if *age* = 20 is specified, the tuples must be partitioned on *age.)* If range selections such as 15 < *age* < 25 are specified, range partitioning is superic)!' to hash partitioning because qualifying tuples are likely to be clustered together on a few processors. On the other hand, range partitioning can lead to data skew; that is, partitions with widely varying numbers of tuples across partitions or disks. Skew causes processors dealing with large partitions to becorne perfonnance bottlenecks. Hash partitioning has the additional virtue that it keeps data evenly distributed even if the data grows and shrinks over tirne.

To reduce skew in range partitioning, the luain question is how to choose the ranges by which tuples are distributed. ()ne effective approach is to take sarnpIes fronl each processor, collect and sort all sarnples, and divide the sorted set of samples into equally sized subsets. If tuples are to be partitioned on *age,* the *age* ranges of the sarnpled subsets of tuples can be used as the basis for redistributing the entire relation.

22.3.2 Parallelizing Sequential Operator Evaluation Code

An elegant software architectnre for parallel DBMSs enables us to readily parallelize existing code for sequentially evaluating a relational oI>crator. The basic idea is to use parallel da.ta streamrlS. Streams (frorn different disks or

the output of other operators) are Inerged as needed to provide the inputs for a relational operator, and the output of an operator is split as needed to parallelize subsequent processing.

A. parallel evaluation plan consists of a dataflow network of relational, luerge, and split operators. 1'he rnerge and split operators should be able to buffer SOlne data and should be able to halt the operators producing their input data. They can then regulate the speed of the execution according to the execution speed of the operator that conSUlues their output.

As we will see, obtaining good parallel versions of algorithllls for sequential operator evaluation requires careful consideration; there is no luagic formula for taking sequential code and producing a parallel version. Good use of split and 111erge in a dataflow software architecture, however, can greatly reduce the effort of implementing parallel query evaluation algorithms, as we illustrate in Section 22.4.3.

22.4 PARALLELIZING INDIVIDUAL OPERATIONS

This section shows how various operations can be implemented in parallel in a shared-nothing architecture. We assurne that each relation is horizontally partitioned across several disks, although this partitioning mayor may not be appropriate for a given query. The evaluation of a query must take the initial partitioning criteria into account and repartition if necessary.

22.4.1 Bulk Loading and Scanning

We begin with two simple operations: *scanning* a relation and *loading* a relation. Pages can be read in parallel while scanning a relation, and the retrieved tuples can then be lnerged, if the relation is partitioned across several disks. More generally, the idea also applies when retrieving all tuples that Incet a selection condition. If hashing or range partitioning is used, selection queries can be answered by going to just those processors that contain relevant tuples.

A sirnilar observation holds for bulk loading. Further, if a relation has associated indexes, any sorting of data entries required for building the indexes during bulk loading can also be done in parallel (see later).

22.4.2 Sorting

A sirnple idea is to let each CPTJ sort the part of the relation that is on its local disk and then rnerge these sorted sets of tuples. The degree of parallelisHl is likely to be lirnited by the rnerging phase.

A better idea is to first redistribute all tuples in the relation using range partitioning. For exarnple, if we want to sort a collection of employee tuples by salary, salary values range froIH 10 to 210, and we have 20 processors, we could send all tuples with salary values in the range 10 to 20 to the first processor, all in the range 21 to 30 to the second processor, and so on. (Prior to the redistribution, while tuples are distributed across the processors, we cannot assurne that they are distributed according to sa1ary ranges.)

Each processor then sorts the tuples assigned to it, using sorne sequential sorting algorithrn. For exaluple, a processor can collect tuples until its lllemory is full, then sort these tuples and write out a run, until all incolning tuples have been written to such sorted runs on the local disk. rrhese runs can then be rnerged to create the sorted version of the set of tuples assigned to this processor. The entire sorted relation can be retrieved by visiting the processors in an order corresponding to the ranges assigned to thenl and sirnply scanning the tuples.

The basic challenge in parallel sorting is to do the range partitioning so that each processor receives roughly the same runnber of tuples; otherwise, a processor that receives a disproportionately large nurnber of tuples to sort becornes a bottleneck and lirnits the scalability of the parallel sort. ()ne good approach to range partitioning is to obtain a sarnple of the entire relation by taking sarnples at each processor that initially contains part of the relation. The (relatively srnall) saruple is sorted and used to identify ranges with equal nUlllbers of tuples. This set of range values, called a splitting vector, is then distributed to all processors and used to range partition the entire relation.

A particularly irnportant application of parallel sorting is sorting the data entries in tree-structured indexes. Sorting data entries can significantly speed up the process of bulk-loading an index.

22.4.3 Joins

In this section, we consider how the join operation can be parallelized.\Ve present the basic idea behind the parallelization and illustrate the use of the rnerge and split operators described in Section 22.:3.2. We focus on parallel hash join, which is widely used, and briefly outline how sort-rnerge join ca,n

be similarly parallelized. ()ther join algorithniS can be parallelized as well, although not as effectively as these two algoritlnns.

Suppose that we want to join two relations, say, A and B, on the *age* attribute. We aSSUIllC that they are initially distribl.lted across several disks in senne way that is not useful for the join operation; that is, the initial partitioning is not based on the join attribute. The basic idea for joining A and B in parallel is to decornpose the join into a collection of k srnnller joins. We can decornpose the join by partitioning both A and B into a collection of k logical buckets or partitions. By using the sarne partitioning function for both A and B, we ensure that the union of the k s1naller joins cOlnputes the join of A and B; this idea is si1nilar to intuition behind the partitioning phase of a sequential hash join, described in Section 14.4.3. Because A and B are initially distributed across several processors, the pa,rtitioning step itself can be done in parallel at these processors. At each processor, all local tuples are retrieved and hashed into one of k partitions, with the salIie hash function used at all sites, of course.

Alternatively, we can partition A and B by dividing the range of the join attribute *age* into k disjoint subranges and placing A and B tuples into partitions according to the subrange to which their *age* values belong. For exanlple, suppose that \ve have 10 processors, the join attribute is *age*, with values froln 0 to IOO. Assurlling uniforrl1 distribution, A and B tuples with $0 \leq age < 10$ go to processor 1, $10 \leq age < 20$ go to processor 2, and so on. This approach is likely to be 1110re susceptible than hash partitioning to data skew (i.e., the number of tuples to be joined can vary widely across partitions), unless the subranges are carefully deterrnined; we do not discuss how good subrange boundaries can be identified.

I-Iaving decided on a partitioning strategy, we can assign each partition to a processor and carry out a local join, using any join algorithrl1 we want, at each processor. In this case, the nUIIlber of partitions k is chosen to be equal to the nUlnber of processors n available for carrying out the join, and during partitioning, each processor sends tuples in the ith partition to processor i. After partitioning, each processor joins the A and$B$ tuples assigned to it. Each join process executes sequential join code and receives input .4. and 13 tuples froro several processors; a rnerge operator lnerges all incorning A tuples, and another merge operator rnerges all incorning B tuples. Depending on 110\V we want tC) distribute the result of the join of A and B, the output of the join process rl1ay be split into several data streallIlS. The network of operators for parallel join is sho\vn in Figure 22.3. To sirnplify the figure, we assurlle that the proc.essors doing the join are distinct frorn the processors that. initially contain tuples of A and B and show only four processors.

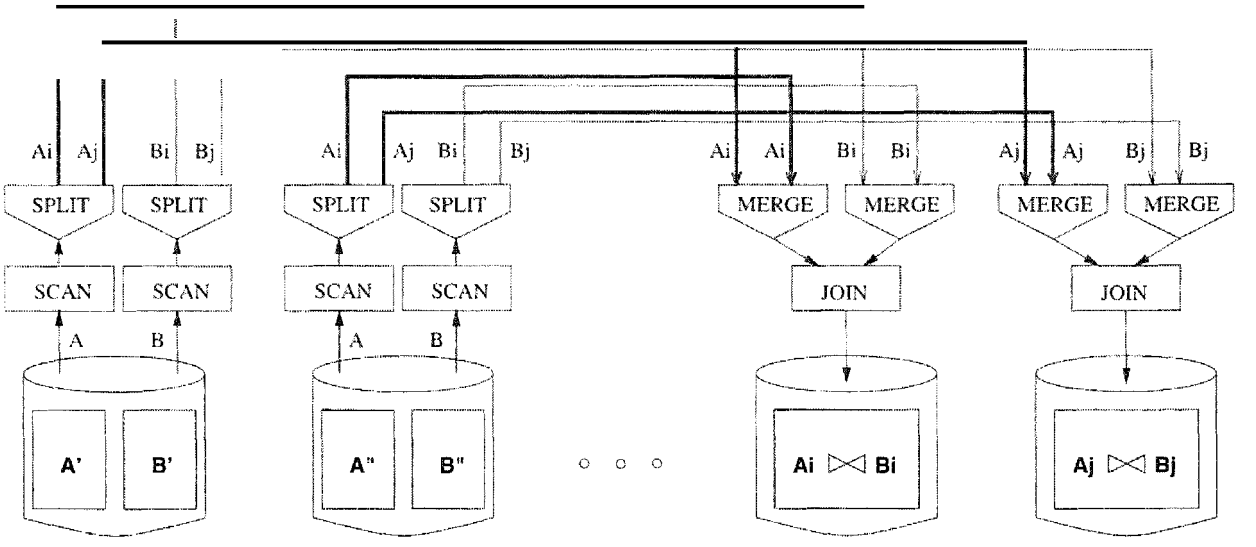


Figure 22.3 Dataflow Network of Operators for Parallel Join

If range partitioning is used, this algorithrn leads to a parallel version of a sort-merge join, with the advantage that the output is available in sorted order. If hash partitioning is used, we obtain a parallel version of a hash join.

Improved Parallel Hash Join

A hash-based refinernent of the approach offers improved perforruanee. The ruain observation is that, if A and B are very large and the nurnber of partitions k is chosen to be equal to the nurnber of processors n, the size of each partition 111ay still be large, leading to a high cost for each local join at the n processors.

An alternative is to execute the srnaller joins $A_i \bowtie B_i$, for i = $1 \ldots k$, one after the other, but\vith each join executed in parallel using all processors. This approa,ch allows us to utilize the total available ruain rueruory at all n processors in each join $A_i \bowtie 13_i$ and is described in rnore detail as follcJ\vs:

1. At each site, apply a hash function hI to partition the A and B tuples at this site into partitions i = $1 \ldots k$. Let A be the srnaller relation. The nurnber of partitions k is chosen such that each partition of A fits into the *aggregate* or cornbined rnernory of all n processors.

2. For $i = 1 \ldots k$, process the join of the ith partitions of A and B. To cornpute $A_i \bowtie B_i$, do the follcnving at every site:
   (a.) i\pply a second hash function $12,2$ to all A_i tuples .to detennine where they should be joined and send tuple t to site $h2(t)$.
   (b) As A_i tuples arrive to be joined, add thcln to an in-rnernory hash. table.

(c) After all A_i tuples have been distributed, apply *h2* to B_i tuples to deterrnine where they should be joined and send tuple *t* to site *h2(t).*

(d) As B_i tuples HJTive to be joined, probe the in-rnernory table of A_i tuples and output result tuples.

The lIse of the second hash function *h2* ensures that tuples are (rIlore or less) uniforrnly distributed across all *n* processors participating in the join. This approach greatly reduces the cost for each of the srnaller joins and therefore reduces the overall join cost. ()bserve that all available processors are fully utilized, even though the srnaller joins are carried out one after the other.

The reader is invited to adapt the network of operators shown in Figure 22.3 to reflect the improved parallel join algorithrn.

22.5 PARALLEL QUERY OPTIMIZATION

In addition to pa.rallelizing individual operations, we can obviously execute different operations in a query in parallel and execute rnultiple queries in parallel. Optirnizing a single query for parallel execution has received rnore attention; systerus typically optirnize queries without regard to other queries that might be executing at the same tilne.

rrwo kinds of interoperatioll parallelisrn can be exploited within a query:

- The result of one operator can be pipelined into another. For example, consider a left-deep plan in which all the joins use index nested loops. The result of the first (i.e., the bottollunost) join is the outer relation tuples for the next join node. As tuples are produced by the first join, they can be used to probe the inner relation in the second join. T'he result of the second join can sirnilarly be pipelined into the next join, and so 011.
- Multiple independent operations can be executed concurrently. For exarnpIe, consider a (bushy) plan in vilhich relations A and B are joined, relations C and D are joined, and the results of these two joins are finally joined. Clearly, the join of A and B can be executed conculTcntly with the join of C and D.

An optirnizer that seeks to parallelize query evaluation has to consider several issues, and we only outline the rnain points. The cost of executing individual operations in paraJlel (e.g., parallel sorting) obviollsly differs frorn executing thern sequentially, and the optirnizer should estimate operation costs accordingly.

Next, the plan that returns answers quickest Inay not be the plan with the least cost. For example, the cost of $A \bowtie B$ plus the cost of C $\bowtie$ D plus the cost of joining their results may be rnore than the cost of the cheapest left-deep plan. However, the time taken is the titne for the 1nore expensive of $A \bowtie B$ and $C \bowtie$ 1) plus the titne to join their results. Tlhis tirHe may be less than the tirne taken by the cheapest left-deep plan. This observation suggests that a parallelizing optirnizer should not restrict itself to left-deep trees and should also consider *bushy* trees, which significantly enlarge the space of plans to be considered.

Finally, a nurnber of pararneters, such as available buffer space and the nUl1.1-bel' of free processors, are known only at run-tirne. rrhis comnlent holds in a rnultiuser environrnent even if only sequential plans are considered; a rnultiuser environrnent is a sirnple instance of interquery parallelisrn.

22.6 INTRODUCTION TO DISTRIBUTED DATABASES

As we observed earlier, data in a distributed database systern is stored across several sites, and each site is typically rnanaged by a DBMS that can run independent of the other sites. The classical view of a distributed database systern is that the systcrn should rnake the irnpact of data distribution transparent. In particular, the following properties are considered desirable:

- Distributed Data Independence: Users should be able to ask queries without specifying where the referenced relations, or copies or fragrnents of the relations, are located. This principle is a natural extension of physical and logical data independence; we discuss it in Section 22.8. Further, queries that span rnultiple sites should be optirnized systcrnatically in a cos1,-based rnanner, taking into account COllllnunication costs and differences in local cornpntation costs. We discuss distributed query optirniza-Lion in Section 22.10.
- Distributed Transaction Atolnicity: Users should be able to write transactions that access and update data at several sites just as they would write transactions over purely local data. In particular, the effects of a transaction across sites should continue to be atornic; that is, all changes persist if the transaction cOllnuits and none persist if it aborts. We discuss this distributed transaction processing in Sections 22.11, 22.13, and 22.14.

AJthough rnost people would agree that these properties are in general clesirable, in certain situations, such as when sites are connected by a slow long-distance network, these properties are not efficiently achievable. Indeed, it has been argued that when sites are globally distributed, these properties are not even desirable. The argurnerlt essentially is that the adrninistrative overhead

of supporting a systern with distributed data independence and transaction atomicity ... in effect, coordinating all activities across all sites to support the view of the whole as a unified collection of data—is prohibitive, over and above DBMS perfo1'rnanc8 considerations.

}(eep these remarks about distributed databases in rnind as we cover the topic in mo1'e detail in the rest of this chapter. There is no real consensus on what the design objectives of distributed databases should be, and the field is evolving in response to users' needs.

22.6.1 Types of Distributed Databases

If data is distributed but all servers run the sarne DBMS software,. we have a homogeneous distributed database system. If different sites run under the control of different DBMSs, essentially autonorllously, and are connected sOlllehow to enable access to data from rnultiple sites, we have a heterogeneous distributed database system, also referred to as a multidatabase system.

The key to building heterogeneous systelTIS is to have well-accepted standards for gateway protocols. A gateway protocol is an API that exposes DBMS functionality to external applications. Examples include ODBC and JDBC (see Section 6.2). By accessing database servers through gateway protocols, their differences (in capability, data fonnat, etc.) are rnasked, and the differences between the different servers in a distributed system are bridged to a large degree.

Gateways are not a panacea, however. They add a layer of processing that can be expensive, and they do not cornpletely mask the differences arllong servers. For example, a server Illay not be capable of providing the services required for distributed transaction rnanagernent (see Sections 22.13 and 22.14), and even if it is capable, standardizing gateway protocols all the way down to this level of interaction poses challenges that have not yet been resolved satisfactorily.

Distributed data rnanagcrnent, in the final analysis, cornes at a significant cost in terulS of performance, software cOlllplexity, and adrninistration difficulty. rrhis observation is especially true of heterogeneous SystCIllS.

22.7 DISTRIBUTED DBMS ARCHITECTURES

Three alternative approaches are used to separat,e functionality across different DBMS-related processes; these alternative distributed])131VI8 architectures are ca,lled *Client-Server*, *Collaborating Server*, and *Middleware*.

22.7.1 Client-Server Systems

A Client-Server systelll has one or mo1'e client processes and one or rnore server processes, and a client process can send a query to anyone server process. Clients are responsible for user-interface issues, and servers rnanage data and execute transactions. Thus, a client process could run on a personal cornputer and send queries to a server running on a 11lainframe.

This architecture has becorne very popular for several reasons. First, it is relatively sinlple to irnplernent due to its clean separation of functionality and because the server is centralized. Second, expensive server rnachines are not underutilized by dealing with lllundane user-interactions, which are now relegated to inexpensive client machines. Third, users can run a graphical user interface that they are familiar with, rather than the (possibly unfalniliar and unfriendly) user interface on the server.

While writing Client-Server applications, it is inlportant to remember the boundary between the client and the server and keep the communication between therll as set-oriented as possible. In particular, opening a cursor and fetching tuples one at a time generates many rnessages and should be avoided. (Even if we fetch several tuples and cache them at the client, rnessages 11lUSt be exchanged when the cursor is advanced to ensure that the current row is locked.) Techniques to exploit client-side caching to reduce comlnunication overhead have been studied extensively, although we do not discuss them further.

22.7.2 Collaborating Server Systems

The (;lient-Server architecture does not allow a single query to span rnultiple servers because the client process would have to be capable of breaking such a query into appropriate subqueries to be executed at different sites and then piecing together the answers to the subqueries. The client process would therefore be quite cOlnplex, and its capabilities would begin to overlap with the server; distinguishing between clients and servers becornes harder. EliIninating this distinction leads us to an alternative to the Client-Server architecture: a Collaborating Server systenl. We can have a collection of database servers, each capable of running tra,nsactions against local data, which cooperatively execute transactions spanning rnultiple servers.

When a server receives a query that requires access to data at other servers, it generates appropriate subqueries to be executed by other servers and puts the results together to COIllpute answers to the original query. Ideally, the decom-

position of the query should be done using cost-based optinlization, taking into account the cost of network COlnnlunication as well as local processing costs.

22.7.3 Middleware Systems

The Middleware architecture is designed to allow a single query to span rnultiple servers, without requiring all database servers to be capable of rnanaging such nlulti-site execution strategies. It is especially attractive when trying to integrate several legacy systerns, whose basic capabilities cannot be extended.

The idea is that we need just one database server capable of rnanaging queries and transactions spanning nlultiple servers; the renlaining servers need to handle only local queries and transactions. We can think of this special server as a layer of software that coordinates the execution of queries and transactions across one or more independent database servers; such software is often called **middleware.** The middleware layer is capable of executing joins and other relational operations on data obtained froln the other servers but, typically, does not itself maintain any data.

22.8 STORING DATA IN A DISTRIBUTED DBMS

In a distributed DBMS, relations are stored across several sites. Accessing a relation stored at a renlote site incurs message-passing costs and, to reduce this overhead, a single relation lnay be *partitioned* or *fragrnented* across several sites, with fragrnents stored at the sites where they are most often accessed or *replicated* at each site where the relation is in high demand.

22.8.1 Fragmentation

Fragrnentation consists of breaking a relation into srnaller relations or fragrnents and storing the fragrnents (instead of the relation itself), possibly at different sites. In horizontal fragmentation, each fragrnent consists of a subset of *rows* of the original relation. In vertical fragluentation, each fragrllent consists of a subset of *columns* of the original relation. IIorizontal and verticaJ fragrnents are illustrated in Figllre 22.4.

Typically, the tuples that belong to a given horizontal fragrnent are identified by a selection query; for exarnple, crnployee tuples Blight be organized into fragments by city, with all enlployees in a given city assigned to the sanie fragrnent. rThe horizontal fragrnent shown in Figure 22.4 corresponds to Chicago. storing fragrncnts in the database site at the corresponding city, we achieve :ality of reference—Chicago data is 1n08t likely to be updated and queried

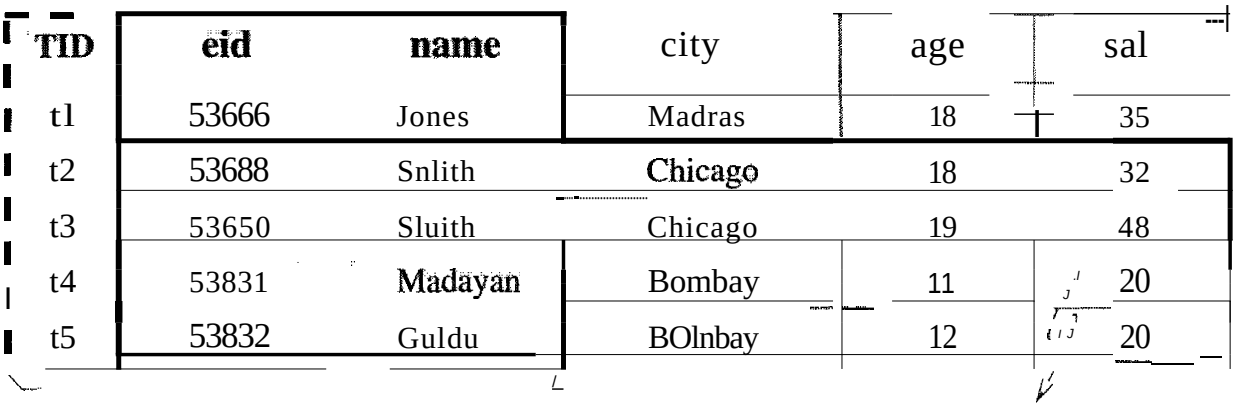


Figure 22.4 Horizontal and Vertical Fragmentation

fronl Chicago, and storing this data in Chicago rnakes it local (and reduces cornrnunication costs) for nlost queries. Sinlilarly, the tuples in a given vertical fragrnent are identified by a projection query. The vertical fragrnent in the figure results frorn projection on the first two columns of the ernployees relation.

When a relation is fragrnented, we lllust be able to recover the original relation fronl the fragrnents:

- Horizontal Fragmentation: The union of the horizontal fragments rnust be equal to the original relation. Fragrnents are usually also required to be disjoint.
- Vertical Fragrnentation: 'The collection of vertical fragrnents should be a lossless-join deccnnposition, as per the definition in Chapter 19.

To ensure that a vertical fragrnentation is lossless-join, systeuls often assign a unique tuple ieI to each tuple in the original relation, as shown in Figure 22.4, and attach this id to the projection of the tuple in each fragrnent. If we think of the original relation as containing an addit.iC)11al tuple-id field that is a key, this field is added to each vertical fragrnent. Such a decol11position is guaranteed to be lossless-join.

In general, a relation can be (horizontally or vertically) fragrnented, a.nd each resulting fragrnent can be further fragnlented. For sirnplicity of exposition, in the rest of this chapter, we assume that fragrnents are not recursively partitioned in this rnanner.

22.8.2 Replication

Replication Incaus that we store several copies of a relation or relation fragrnent. An entire relation can be replicated at one or rnore sites. Sirnila,rly, one or 1110re fragrncnts of a relation can be replicated at other sites. For example, if a relationR is fragrnented into $R1$, $R2$, and $R3$, there nlight be just one copy of $R1$, whereas *R2* is replicated at two other sites and $R3$ is replicated at all sites.

The rnotivation for replication is twofold:

- Increased Availability of Data: If a site that contains a replica goes down, we can find the sarne data at other sites. Sirnilarly, if local copies of rerllote relations are available, we are less vulnerable to failure of COllnnunication links.
- Faster Query Evaluation: Queries can execute faster by using a local copy of a relation instead of going to a rernote site.

The two kinds of replication, called *synchronous* and *asynchronous* replication, differ prirnarily in how replicas are kept current when the relation is rnodified (see Section 22.11).

22.9 DISTRIBUTED CATALOG MANAGEMENT

Keeping track of data distributed across several sites can get cornplicated. We rnust keep track of how relations are fragrnented and replicated------that is, how relation fragrnents are distributed across several sites and where copies of fragrnents are stored—in addition to the lIsuaJ seherna, authorization, and statistical inforrnation.

22.9.1 Naming Objects

If a relation is fragruented and replicated, we rnust be able to uniquely identify each replica of each fragnlent. Generating such unique narnes requires sorne care. If we use a global narne-server to assign globally unique narnes, local autonomy is cornprornised; we 'want (users at) each site to be able to assign names to local objects without reference to names systernwide.

The usual solution. to the naTning problenl is to use names consisting of several fields. 1;'01' example, we could have:

- A *local name* field, 'which is the name assigned locally at the site\vhere the relation is created. T\vo objects at different sites could have the saIne local name, but two objects at a given site cannot have the saIne local narne.
- A *birth site* field, which identifies the site where the relation was created, and where il1fofrnation is ruaintained about all fragruents and replicas of the rela.tion.

These two fields identify a relation uniquely; we call the cornbination a global relation nanle. To identify a replica (of a relation or a relation fragnlent),we take the global relation narne and add a *replica-id* field; we call the cornbination a global replica narrle.

22.9.2 Catalog Structure

A centralized systern catalog can be used but is vulnerable to failure of the site containing the catalog. An alternative is to rnaintain a copy of a global system catalog,which describes all the data at every site. Although this approach is not vulnerable to a single-site failure, it comprornises site autonorny, just like the first solution, because every change to a local catalog rnust now be broadcast to all sites.

A better approach, which preserves local autonoruy and is not vulnerable to a single-site failure, was developed in the R* distributed database project, which was a successor to the Systerl1 R project at IBlV!. Each site ruaintains a local catalog that describes all copies of data stored at that site. In addition, the catalog at the birth site for a relation is responsible for keeping track of where replicas of the relation (in general, of fragnlents of the relation) are stored. In particular, a precise description of each replica's contents—a list of colurl1ns for a vertical fragrnent or a selection condition for a horizontal fragruentis stored in the birth site catalog. Whenever a new replica is created or a replica is rnoved across sites, the inforrnation in the birth site catalog for the relation HUlst be updated.

To locate a relation, the catalog; at its birth site lnust be looked up. This catalog inforrnation can be ca.,ched at other sites for quicker access, but the cached inforrnation Inay becolue out of date if, for cxarnple, a fragrnent is rnoved. We would discover that the locally cached inforrnation is out of date when \ve use it to access the relation, and at that point, we rll11st update the cache b.y looking up the catalog at the birth site of the relation. (The birth site of a relation is recorded in each local cache that describes the relation, and the birth site never changes, even if the relation is rnoved.)

22.9.3 Distributed Data Independence

Distributed data independence lueans that users should be able to write queries \vithout regard to ho\v a relation is fragrnented or replicated; it is the responsibility of the DBMS to cornpute the relation as needed (by locating suitable copies of fragrnents, joining the vertical fragrnents, and taking the union of horizontal fragn1cnts).

In particular, this property irnplies that users should not have to specify the full nalne for the data objects accessed while evaluating a query. Let us see how users can be enabled to access relations without considering how the relations are distributed. The *local narne* of a relation in the systeln catalog (Section 22.9.1) is really a c01l1bination of a *user narne* and a user-defined *relation narne.* 'Users can give whatever names they wish to their relations, without regard to the relations created by other users. When a user writes a prograrn or SQL statelnent that refers to a relation, he or she sirnply uses the relation narne. The DBMS adds the user narne to the relation narne to get a local narne, then adds the user's site-id as the (default) birth site to obtain a global relation narne. By looking up the global relation narne--in the local catalog if it is cached there or in the catalog at the birth site-·the DBMS can locate replicas of the relation.

A user Il1ay want to create objects at several sites or refer to relations created by other users. To do this, a user can create a synonym for a global relation narne' llsing an SQL-style cOl1unand (although such a corl1rnand is not currently part of the SQL:1999 standard) and subsequently refer to the relation using the synonyrn. For each user known at a site, the DBMS maintains a table of synonynls as part of the systern catalog at that site and uses this table to find the global relation narne. Note that a user's prograrl1 runs unchanged even if replicas of the relation are rlloved, because the global relation narne IS never changed until the relation itself is destroyed.

lJsers rnay want to run queries against specific replicas, especially if asynchronous replication is used. To support this, the synonyrn Inechanisrn can be adapted to also allo\v users to create synon.yrl1S for global replica, names.

22.10 DISTRIBUTED QUERY PROCESSING

We first discuss the issues involved in evaluating relational algebra operations in a distrilnlted database through exalnples and then outline distributed query optiInization. Consider the following tvo relations:

Sailors(*sid:* integer, *sname:* string, *rating:* integer, *age:* real)
Reserves(*sid:* integer, *bid:* integer, *day:* date, *rname:* string)

As in Chapter 14, assurne that each tuple of R.eserves is 40 bytes long, tha.t a page can hold 100 Reserves tuples, and that we have 1000 pages of such tuples. Sirnilarly, assurne that each tuple of Sailors is 50 bytes long, that a page can hold 80 Sailors tuples, and that we have 500 pages of such tuples.

To estimate the cost of an evaluation strategy, in addition to counting the nurnber of page IjC)s, \ve Blust count the nurnher of pages sent frorH one site to another because corllrIlunication costs are a significant cornponent of overall cost in a distributed database. We rnust also change our cost rnodel to count the cost of shipping the result tuples to the site where the query is posed frOIn the site where the result is assernbled! In this chapter, we denote the time taken to read one page from disk (or to write one page to disk) as t_d and the tiIne taken to ship one page (from any site to another site) as t_s.

22.10.1 Nonjoin Queries in a Distributed DBMS

Even sirnple operations such as scanning a relation, selection, and projection are affected by fragmentation and replication. Consider the following query:

```
SELECT  S.age
FROM    Sailors S
WHERE   S.rating > 3 AND S.rating < 7
```

Suppose that the Sailors relation is horizontally fragruented, with all tuples having a rating less tha,n 5 at Shanghai and all tuples having a rating greater than 5 at rrokyo.

TheDBl\/IS nn.lst answer this query by evaluating it at both sites and taking the union of the ans\vers. If the SELECT clause contained AVG *(S. age),* COIn- bining the answers could not be done by sirnply taking the union····-····theDBMS rnust cornpute the SUIn and count of *age* values at the two sites and use this infonna,tion to cornpute the average age of all sailors.

If the WHERE clause contained just the condition 5'.*rating* > *6*, on the other ha,ud, the I)BI\/IS should recognize that this query could be answered by just executing Ht-at Tnkyo.

As another example, suppose that the Sailors relation, were vertically frag- mented, with the *sid* and *rating* fields at Sha.nghai a,nd the *sname* and *age* fields at rrokyo. No field is stored at both sites. This vertical fragmentation

would therefore be a lossy decornposition, except that a field containing the id of the corresponding Sailors tuple is included by the DBMS in both fraglnents! Now, the DBMS has to reconstruct the Sailors relation by joining the t\VO fraglnents on the COHIIIIon tuple-id field and execute the query over this reconstructed relation.

Finally, suppose that the entire Sailors relation were stored at both Shanghai and Tokyo. We could answer any of the previous queries by executing it at either Sha,nghai or Tokyo. Where should the query be executed? This depends on the cost of shipping the answer to the query site (which rnay be Shanghai, Tokyo, or SOlllle other site) as well as the cost of executing the query at Shanghai and at Tokyo—···the local processing costs lnay differ depending on what indexes are available on Sailors at the two sites, for exaluple.

22.10.2 Joins in a Distributed DBMS

Joins of relations at different sites can be very expensive, and we now consider the evaluation options that IllUst be considered in a distributed environrnent. Suppose that the Sailors relation were stored at London, and the Reserves relation were stored at Paris. We consider the cost of various strategies for cOlnputing *Sailor'S* $\bowtie$ *Reserves.*

Fetch As Needed

We could do a page-oriented nested loops join in Loudon with Sailors as the outer, and for each Sailors page, fetch all Reserves pages frorn Paris. If we cache the fetched Reserves pages in London until the join is complete, pages are fetched only once, but aSSU111e that H,eservcs pages are not cached, just to see how bad things can get. (The situation can get rnuch worse if we use a tuple-oriented nested loops join!)

rrhe cost is $500t_d$ to scan Sailors plus, for each Sailors page, the cost of seallning and shipping all of Reserves, which is 1000(td + *ts).* The total cost is therefore *500td* + 500,OOO(td + t_s).

In addition, if the query was not sllbrnittccl at the London site, we rnust add the cost of shipping the result to the query site; this cost depends on the size of the result. Because *sid* is a key for Sailors, the nurnber of tuples in the result is 100,000 (the rnunber ()f tuples in Reserves) and each tuple is 40 + 50 = 90 bytes long; thus 4000/90 = 44 result tuples fit on a page, and the result size is 100,000/44=2273 pages. The cost of shipping the answer to another site, if necessary, is 2273 t_s. In tIle rest of this section, we assume that the query is

posed at the site where the result is computed; if not, the cost of shipping the result to the query site Blust be added to the cost.

In this exarnple, observe that, if the query site is not London or Paris, the cost of shipping the result is greater than the cost of shipping both Sailors and Reserves to the query site! Therefore, it would be cheaper to ship both relations to the query site and COlllpute the join there.

Alternatively, we could do an index nested loops join in London, fetching all Inatching Reserves tuples for each Sailors tuple. Suppose we have an uncIustered hash index on the *sid* colurnn of Ileserves. Because there are 100,000 Ileserves tuples and 40,000 Sailors tuples, each sailor has on average 2.5 reservations. The cost of finding the 2.5 Ileservations tuples that lllatch a given Sailors tuple is $(1.2 + 2.5)td'$ assluning 1.2 l/Os to locate the appropriate bucket in the index. The total cost is the cost of scanning Sailors plus the cost of finding and fetching nlatching Reserves tuples for each Sailors tuple, $500td + 40,000(3.7td + 2.5t_s)'$

Both algorithIns fetch required Reserves tuples from a remote site as needed. Clearly, this is not a good idea; the cost of shipping tuples dominates the total cost even for a fast network.

Ship to One Site

We can ship Sailors from London to Paris and carry out the join there, ship Reserves to London and carry out the join there, or ship both to the site where the query was posed and cornpute the join there. Note again that the query could have been posed in London, Paris, or perhaps a third site, say, Tirnbuktu!

1'he cost of scanning and shipping Sailors, saving it at Paris, then doing the join at Paris is $500(2td + t_s) + 4500t_d$, assurning that the version of the sort-rnerge join described in Section 14.10 is used and we have an adequate nurnber of buffer pages. In the rest of this section we aSSUInc that sort-Inerge join is the join rnethod used when both relations are at the saIne site.

The cost of shipping Reserves and doing the join at London is $1000(2t(1 + t_s) + 4500td$·

Senlijoins andBloomjoins

Consider the strategy of shipping Reserves to Londo.l1 and cornputing the join at London. Some tuples in (the current instance of)H,cserves do not join with an.y tuple in (the current instance of) Sailors. If we could somehow identify

Reserves tuples that are guaranteed not to join with any Sailors tuples, we could avoid shipping thern.

Two techniques, *Semijoin* and *Bloomjoin*, have been proposed for reducing the number of lleserves tuples to be shipped. The first technique is called Semijoin. The idea is to proceed in three steps:

1. At London, cornpute the projection of Sailors onto the join colurnns (in this case just the *sid* field) and ship this projection to Paris.
2. At Paris, cornpute the natural join of the projection received frorn the first site with the R,eserves relation. The result of this join is called the reduction of R,eserves with respect to Sailors. Clearly, only those Reserves tuples in the reduction will join with tuples in the Sailors relation. Therefore, ship the reduction of Reserves to London, rather than the entire Reserves relation.
3. At London, cornpute the join of the reduction of R,eserves with Sailors.

Let us compute the cost of using this technique for our example join query. Suppose we have a straightforward irnplernentation of projection based on first scanning Sailors and creating a telnporary relation with tuples that have only an *sid* field, then sorting the temporary and scanning the sorted ternporary to eliminate duplicates. If we assurne that the size of the *sid* field is 10 bytes, the cost of projection is $500t_d$ for scanning Sailors, plus 100td for creating the ternporary, plus $400t_d$ for sorting it (in two passes), plus IOOtc! for the final scan, plus $100t_d$ for writing the result into another tcrnporary relation; a total of $1200t_d$. (Because *sid* is a key, no duplicates need be elirninated; if the optiInizer is good enough to recognize this, the cost of projection is just (500 + 100)td.)

The cost of cornputing the projection and shipping it to Paris is therefore $1200ld + 100t_s$. The cost of c(nnputing the reduction of R.eserv8s is $3 . (100 + 10(0) = 3300t_d$, assurning that sort-rnerge join is used. (The cost does not reflect that the projection of Sailors is already sorted; the cost would decrease slightly if the refined sort-Inerge join exploited this.)

What is the size of the reduction? If every sailor holds at least one reservation, the reduction includes every tuple of R,eserves! The effort invested in shipping the projection and reducing Reserves is a total waste. Indeed, because of this observation, we note that Sernijoin is especially useful in conjunction with a selectiol1 on one of the relations. For example, if we want to cornpute the join of Sailors tuples with a *rating* greater than 8 with the Reserves rela.tion, the size of the projection on *sid* for tuples that satisfy the selection would be just 20 percent of tIle original projection, that is, 20 pages.

Let us now continue the example join, 'with the assurnption that we have the additional selection on *rating.* (The cost of cornputing the projection of Sailors goes down a bit, the cost of shipping it goes down to $20t_s$, and the cost of the reduction of Reserves also goes down a little, but we ignore these reductions for sirnplicity.) \Ve assurne that; only 20 percent of the Reserves tuples are included in the reduction, thanks to the selection. Hence, the reduction contains 200 pages, and the cost of shipping it is $200t_s$.

Finally, at London, the reduction of I{eserves is joined with Sailors, at a cost of $3 \cdot (200 + 500) = 2100td$. Observe that there are over 6500 page I/Os versus about 200 pages shipped, using this join technique. In contrast, to ship R,eserves to London and do the join there costs $1000t_s$ plus 4500td. With a high-speed. network, the cost of Sernijoin Illay be n10re than the cost of shipping Reserves in its entirety, even though the shipping cost, itself is rnueh less ($200t_s$ versus $1000t_s$).

The second technique, called Bloomjoin, is quite sirnilar. The luain difference is that a bit-vector is shipped in the first step, instead of the projection of Sailors. A bit-vector of (SOIIIC chosen) size k is cOlnputed by hashing each tuple of Sailors into the range 0 to k - I and setting bit i to I if seHne tuple hashes to i, and 0 otherwise. In the second step, the reduction of Reserves is cOlnputed by hashing each tuple of Reserves (using the *sid* field) into the range 0 to k --1, using the sanle hash function used to construct the bit-vector and discarding tuples whose hash value i corresponds to a 0 bit. Because no Sailors tuples hash to such an i, no Sailors tuple can join with any R,eserves tuple that is not in the reduction.

The costs of shipping a bit-vector and reducing R,eserves using the vector are less than the corresponding costs in Sernijoin. ()n the other hand, the size of the reduction of Reserves is likely to be larger than in Sernijoin; so, the costs of shipping the reduction and joining it 'with Sailors are likely to be higher.

Let us estirnate the cost of this approach. rrhe cost of cornputing the bit-vector is essentially the cost of scanning Sailors, \vhich is *500td.* The cost of sending the bit-vector depends on the size we choose for the bit-vector, 'which is certainly sInaJler than the size of the projection; we take this cost to be $201:_8$, for concreteness. The cost of reducing Reserves is just the cost of scanning Reserves, $1000t_d$. T'he size of the reduction of Reserves is likely to be about the saIne as or a little larger than the size of the reduction in tIle Scrnijoin approach; instea,d of 200, we will take tllis size to be 220 pages. (We assume that the selection on Sailors is included, to pennit a direct cOInparison \vith the cost of Scrnijoin.) 'I'he cost of shipping the reduction is therefore $220t_s$' The cost of the final join at London is $3 . (500 + 220) = 2160td$.

rrhus, in cornparison to Semijoin, the shipping cost of this approach is about the same, although it could be higher if the bit-vector were not as selective as the projection of Sailors in terms of reducing Reserves. Typically, though, the reduction of Reserves is no l110re than 10 to 20 percent larger than the size of the reduction in SClnijoin. In exchange for this slightly higher shipping cost, Bloornjoin achieves a significantly lower processing cost: less than $3700t_d$ versus rnore than 6500td for SClnijoin. Indeed, Bloornjoin has a lower *I/C)* cost and a lower shipping cost than the strategy of shipping all of R,eserves to London! These nurnbers inclicatewhy Bloollljoin is an attractive distributed join rnethod; but the sensitivity of the rnethod to the effectiveness of bit-vector hashing (in reducing Reserves) should be kept in rnind.

22.10.3 Cost-Based Query Optimization

We have seen how data distribution can affect the inlplernentation of individual operations, such as selection, projection, aggregation, and join. In general, of course, a query involves several operations, and optirnizing queries in a distributed database poses the following additional challenges:

- CornrIlunication costs lllUSt be considered. If we have several copies of a relation, we HlllSt also decide which copy to use.
- If individual sites are run under the control of different DBl\JlSs, the autonolny of each site HUlst be respected while doing global query planning.

Query optiInization proceeds essentially as in a centralized DBMS, as described in Chapter 12, with inforrnation about relations at rernote sites obtained fron1 the systeln catalogs. ()f course, there are nlore alternative lllethods to consider for each operation (e.g., consider the new options for distributed joins), and the cost rnetric ruust account for cornrnunication costs as \vell, but the overall planning process is essentially unchanged if we take the cost rnetric to be the total cost of all operations. (If we consider response tilne, the fact that certain subqueries can be carried out in parallel at different sites \vould require us to change the optirnizer as per the discussion in Section 22.5.)

In the overall plan, local rnanipulatiol1 of relations at the site where they are stored (to corllpute an interrnediate relation to be shipped elsewhere) is encapsulated into a *suggested* local plarl. The overall plan includes several such local plans, \vhichwe can think of as subqueries executing at different sites. \:Vhile generating the global plan, the suggested local plans provide realistic cost estimates for the cornputatioll of the interrnediate relations; the suggested local plans are constructed by the optirnizer rnainly to provide these local cost estiIuates. A site is free to ignore the local plan suggested to it if it is able to find a cheaper plan by llsing more current infonnation in the local catalogs. Thus,

site autonomy is respected in the optimization and evaluation of distributed quenes.

22.11 UPDATING DISTRIBUTED DATA

The classical view of a distributed DBMS is that it should behave just like a centralized DBMS froul the point of view of a user; issues arising froln distribution of data should be transparent to the user, although, of course, they rnu8t be addressed at the irnplernentation level.

With respect to queries, this view of a distributed DBIVIS Ineans that users should be able to ask queries \vithout worrying about how and where relations are stored; we have already seen the irnplications of this requirernent on query evaluation.

With respect to updates, this view rneans that transactions should continue to be atornic actions, regardless of data fragrnentation and replication. In particular, all copies of a rnodified relation must be updated before the rnodifying transaction cornn1its. We refer to replication with this sernantics as synchronous replication; before an update transaction cOHllnits, it synchronizes all copies of rnodified data.

An alternative approach to replication, called asynchronous replication, has corne to be widely usee1 in eornrnercial distributed DBlVISs. Copies of a rnodified relation are updated only periodically in this approach, and a transaction that reads different copies of the sarne relation rnay see different values. T'hus, asynchronous replication cornprolnises distributed data independence, but it can be iInplernented 1110re efficiently than synchronous replication.

22.11.1 Synchronous Replication

There are two basic techniques for ensuring that transactions see the same value regardless of\vhich copy of an object they access. In the first technique, called voting, a transaction Inust write H' lnajority of copies to rnodify an ol)ject and read at least enough copies to rnake sure that one of the copies is current. For exanlple, if there are 10 copies and 7 copies are written by update transactions, then at least 4 copies rnust be read. Eac:h copy has a version nurnber, and the copy with the highest version rllunber is current. This technique is not attra,ctive in rnost situations because reading an ol)ject reqllires reading rnultiple copies; in rnost applications, objects are read rnuch 1n01'e frequently than they are updated, and efficient performance on reads is very irnportant.

In the second technique, called read-any write-all, to read an object, a traJ1S-action can read anyone copy, but to write an object, it Inust write all copies. Reads are fast, especially if we have a local copy, but 'writes are slower, relative to the first technique. This technique is attractive when reads are rnuch rnore frequent than writes, and it is usually adopted for irnplernenting synchronous replication.

22.11.2 Asynchronous Replication

Synchronous replication COines at a significant cost. Before an update transaction can corn1'nit, it rnust obtain exclusive locks on all copies—assuming that the read-any write-all technique is used· ··ofrnodified data. The transaction Inay have to send lock requests to rernote sites and wait for the locks to be granted, and during this potentially long period, it continues to hold all its other locks. If sites or connnunication links fail, the transaction cannot cOInrnit until all the sites at which it has rnodified data recover and are reachable. Finally, even if locks are obtained readily and there are no failures, connnitting a transaction requires several additional rnessages to be sent as part of a *commit protocol* (Section 22.14.1).

For these reasons, synchronous replication is undesirable or even unachievable in 111any situations. Asynchronous replication is gaining in popularity, even though it allows different copies of the saIne object to have different values for short periods of tinlC. This situation violates the principle of distributed data independence; users 11Ulst be aware of which copy they are accessing, recognize that copies are brought up-to-date only periodically, and live with this reduced level of data consistency. Nonetheless, this seeIns to be a practical C0l11pr0l11ise that is acceptable in rnany situations.

Primary Site versus **Peer-to-Peer** Replication

A.synchronous replication C01nes in two flavors. In primary site asynchronous replication, one copy of a relation is designated the primary or luaster copy. H.eplicas of the entire relation or fragrnents of the relation can be created at other sites; these are secondary copies, and unlike tIle primary copy, they cannot be updated. A conUIlon InecllallislIl for setting up primary and secondary copies is that users first register or publish the relation at the primary site and subsequently subscribe to a fragment of a registered relation fron1 another (secondary) site.

In peer-to-peer asynchronous replication, 111()re than one copy (although perhaps rIot all) can be designated as updatable, that is, a 1'naste1' copy. In addition to propagating changes, a conflict resolution strategy must be used to deal

with conflicting changes Inade at different sites. For example, .Joe's age may be changed to 35 at one site and to 38 at another. Which value is 'correct'? Many luore subtle kinds of conflicts can arise in peer-to-peer replication, and in general peer-to-peer replication leads to ad hoc conflict resolution. Some special situations in which peer-to-peer replication does not lead to conflicts arise quite often, and in such situations peer-to-peer replication is best utilized. For example:

- Each 1naster is allo\ved to update only a fragrnent (typically a horizontal frag1nent) of the relation, and any two fragrnents updatable by different !llasters are disjoint. For example, it rIlay be that salaries of Gerrnan erllployees are updated only in Frankfurt, and salaries of Indian ernployees are updated only in 1\1adras, even though the entire relation is stored at both Frankfurt and Madras.
- Updating rights are held by only one rnaster at a tillIe. For example, one site is designated a *backup* to another site. Changes at the Iuaster site are propagated to other sites and updates are not allowed at other sites (including the backup). But, if the Iuaster site fails, the backup site takes over and updates are now perrnitted at (only) the backup site.

We will not discuss peer-to-peer replication further.

Implementing Primary Site Asynchronous Replication

The Inain issue in irnpler11enting prilnary site replication is deterrnining how changes to the prirnary copy are propagated to the secondary copies. Changes are usually propagated in two steps, called *Capture* and *Apply.* Changes rnade by cOHnnitted transactions to the prirnary copy are s()Jnehow identified during the Capture step and subsequently propagated to secondary copies during the Apply step.

In contrast to synchronous replication, a transacti.on that rnodifies a replicated relation directly locks and changes only the prirnary copy. It is typically C0111rnitted long before the Apply step is carried out. Systcrnsvary considerably in their ilnplernentation of these steps. We present an overview of some of the alternatives.

Capture

rrlle Capture step is implemented using one of two approaches. In log-based Capture. the log luainta,inecl for recovery purposes is used to generate a record of updates. Basically, when the log tail is written to stable storage, all log

records that affect replicated relations are also written to a separate change data table (eDT). Since the transaction that generated the update log record may still he active when the record is\vritten to the CDT, it may subsequently abort. lJpdate log records written by transactions that subsequently abort 1l1USt l)e rcrnoved frorl1 the eDT to obtain a strearll of updates due (only) to conlln,itted transactions. This streanl can be obtained as part of the Capture step or subsequently in the Apply step if conunit log records are added to the eDT; for concreteness, we aSSUlne that the cornruitted update strealll is obtained as part of the Capture step and that the CDT sent to the Apply step contains only update log records of corl1ruitted transactions.

In procedural Capture, a procedure autornatically invoked by the DBMS or an application progra,lIl initiates the Capture process, which consists typically of taking a snapshot of the prirnary copy. A snapshot is just a copy of the relation as it existed at sorne instant in tirne. (A procedure that is autoluatically invoked by the DBMS, such as the one that initiates Capture, is called a *trigger*. We covered triggers in Chapter 5.)

Log-based Capture has a s111aller overhead than procedural Capture and, because it is driven by changes to the data, results in a slua11er delay between the tirne the prirnary copy is changed and the ti111e that the change is propagated to the secondary copies. (Of course, this delay also depends on ho\v the Apply step is implelnented.) In particular, only changes are propagated, and related changes (e.g., updates to two tables with a referential integrity constraint between thern) are propagated together. The disadvantage is that ilnpleluenting log-based Capture requires a detailed understanding of the structure of the log, which is quite systern specific. Therefore, a· vendor cannot easily implement a log-based Capture rnechanisrn that will capture changes rnade to data in another vendor's DBMS.

Apply

fI'he Apply step takes the changes collected by the Capture step, which are in the CDT table or a snapshot, and propagates 1,h81n to the secondary copies. This can be done by having the prirnary site continuously send the CDT or periodically requesting (the latest portion of) the CDT or a snapshot frorH the prirnary site. Typically, each secondary site runs a copy of the J\pply process and 'pulls' the changes in the eDT fronl the prirnary site using periodic requests. The interval between such requests can be controlled by a timer or a user's application prograrl1. ()nce the changes are avail(1)le at the secondary site, they can be applied directly to the replica.

In sorne systems, the replica Heed not be just a frag1Ilent of the original relation it can be a view defined using SQL, and the replication rnechanisrn is sufficiently sophisticated to 11laintain such a view at a reillote site incrementally (by reevaluating only the part of the vie\v affected by changes recorded in the CI)T).

Log-based Capture in conjunction with continuous Apply rninirnizes the delay in propagating changes. It is the best cor11bination in situations where the primary and secondary copies are both used as part of an operational DBMS and replicas must be as closely synchronized with the prinlary copy as possible. Log-based Capture with continuous Apply is essentially a less expensive substitute for synchronous replication. Procedural Capture and application-driven Apply offer the 11l0St flexibility in processing source data and changes before altering the replica; this flexibility is often useful in data warehousing applications where the ability to 'clean' and filter the retrieved data is 1110re important than the currency of the replica.

Data Warehousing: An Example of Replication

Cornplex decision support queries that look at data from Illultiple sites are becoming very inlportant. The paradigrn of executing queries that span r11ultiple sites is sirnply inadequate for perfornlance reasons. One way to provide such complex query support over data froln rllultiple sources is to create a copy of all the data at SaIne one location and use the copy rather than going to the individual sources. Such a copied collection of data is called a data warehouse. Specialized systelIls for building, rnaintaining, and querying data warehouses have becolne irnportant tools in the rnarketplace.

Data warehouses can be seen as one instance of asynchronous replication, in 'which copies are updated relatively infrequently. When we talk of replication, we typically rIlCal1 copies Inaintained under the control of a single DBMS, \vhereaswith data \varehousing, the original data rnay be on different software platforrns (including database systems and as file systerIls) and even l)clong to different organizations. This distinction, 110\VeVer, is likely to becoine blurred as vendors adopt luore 'open' strategies to replication. For exarnple, sorne products already support the maintenance of replicas of relations stored in one vendor's DBMS in al10ther vendor's DBMS.

We 110te that data warehousing involves rnore than just replication. We discuss other aspects of data warehousing in Chapter 2.5.

22.12 DISTRIBUTED TRANSACTIONS

In a distributed DBMS, a given transaction is subrnitted at SOlne one site, but it can access data at other sites as well. In this chapter we refer to the activity of a transaction at a given site as a subtransaction. When a transaction is subrnitted at S0111e site, the transaction rnanager at that site breaks it up into a collection of one or rnoro subtransactions that execute at different sites, subrnits theln to transaction rnanagers at the other sites, and coordinates their activity.

We now consider aspects of concurrency control and recovery that require additional attention because of data distribution. As we saw in Chapter 16, there are many concurrency control protocols; in this chapter, for concreteness, we assurne that Strict 2PL with deadlock detection is used. We discuss the following issues in subsequent sections:

- Distributed Concurrency Control: How can locks for objects stored across several sites be managed? How can deadlocks be detected in a distributed database?
- Distributed Recovery: Transaction atomicity lllUSt be ensured-·---when a transaction commits, all its actions, across all the sites at which it executes, rnust persist. Si111ilarly, when a transaction aborts, none of its actions must be allowed to persist.

22.13 DIS"fRIBUTED CONCURRENCY CONTROL

In Section 22.11.1, we described tvo techniques for irnplernenting synchronous replication, and in Section 22.11.2, we discussed various techniques for irllplernenting asynchronous replication. rrhe choice of technique deterrnines *which* objects are to be locked. *When* locks are obtained and released is deterrnined by the concurrency control protocol.vVe now consider how lock and unlock requests are implemented in a distributed environllnent.

Lock rnanagernent can be distributed across sites in rnanyways:

- Centraliz,ed: A single site is in charge of handling lock and unlock requests for all objects.
- Priulary Copy: ()ne copy of each object is designated the primary copy. All requests to lock or unlock a copy of this object are handled by the lock rnanager at the site where the prirnary copy is stored, regardless of where the copy itself is stored.

- Fully Distributed: R,equests to lock or unlock a copy of an object stored at a site are handled by the lock lnanager at the site where the copy is stored.

The centralized schelne is vulnerable to failure of the single site that controls locking. The prirnary copy scherne avoids this problern, but in general, reading a,n object requires cornrnunicatiollwith t\VO sites: the site where the prirnary copy resides and the site where the copy to be read resides. This problern is avoided in the fully distributed 8che1nc, because locking is done at the site where the copy to be read resides. However, while writing, locks rnust be set at all sites where copies are rnoclified in the fully distributed schclne, whereas locks need be set only at one site in the other two schernes.

Clearly, the fully distributed locking scherne is the 1110st attractive schelne if reads are much more frequent than writes, as is usually the case.

22.13.1 Distributed Deadlock

One issue that requires special attention when using either priluary copy or fully distributed locking is deadlock detection. (Of course, a deadlock prevention scherne can be used instead, but we focus on deadlock detection, which is widely used.) As in a centralized DBMS, deadlocks rnust be detected and resolved (by aborting sorne deadlocked transaction).

Each site rnaintains a local waits-for graph, and a cycle in a local graph indicates a, deadlock. lIowever, there can be a deadlock even if no local graph contains a cycle. For exarnple, suppose that two sites, A and B, both contain copies of objects 01 and 02, and that the read-any write-all technique is used. $T1$, which wants to read ()1 and write 02, obtains an S lock on 01 and an X lock on 02 at Site A, then requests an X lock on 02 at Site B. *T2,* which \vants to read 02 and write 01, rneanwhilc, obtains an S lock on 02 and an X lock on 01 at Site B, then requests an X lock on ()1 at Site A. As Figure 22.5 illustrates, $T2$ is waiting for Tl at Site A. and Tl is waiting for $T2$ at Site 13; thus, we have a deadlock, \vhich neither site can detect based solely on its local waits-for graph.

To detect such deadlocks, a distributed deadlock detection algoritlun rnust be used. We describe three such algoritluns.

The first algorithrn,\vhich is centralized, consists of periodically sending all 10-cal waits-for graphs to one site that is responsible for global deadlock detection. At this site, the global waits-for graph is generated by cOlubinin.g all the local graphs; the set of nodes is the union of nodes in the local graphs, and there is

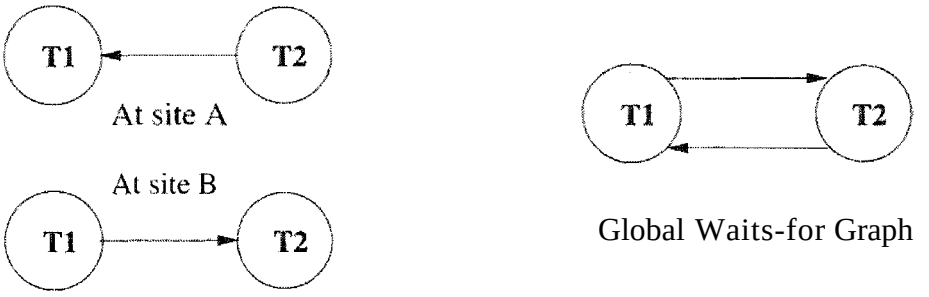


Figure 22.5 Distributed Deadlock

an edge frorn one node to another if there is such an edge in any of the local graphs.

The second algorithrn, which is hierarchical, groups sites into a hierarchy. For instance, sites r11ight be grouped by state, then by country, and finally into a single group that contains all sites. Every node in this hierarchy constructs a waits-for graph that reveals deadlocks involving only sites contained in (the subtree rooted at) this node. All sites periodically (e.g., every 10 seconds) send their local waits-for graph to the site responsible for constructing the waits-for graph for their state. The sites constructing waits-for graphs at the state level periodically (e.g., every minute) send the state waits-for graph to the site constructing the waits-for graph for their country. The sites constructing waits-for graphs at the country level periodically (e.g., every 10 rninutes) send the country waits-for graph to the site constructing the global waits-for graph. This scheme is based on the observation that l110re deadlocks are likely across closely related sites than across unrelated sites, and it puts 1110re effort into detecting deadlocks across related sites. All deadlocks are eventually detected, but a deadlock involving two different countries J.nay take a while to detect.

The third algorithrn is sirllple: If a transaction waits longer than SOIne chosen tinle-out interval, it is aborted. Although this algorithrll rnay cause rnany unnecessary restarts, the overhead of deadlock detection is (obviously!) low, and in a heterogeneous distributed database, if the participating sites cannot cooperate to the extent of sha,ring their \va,its-for graphs, it rnay be the only option.

A subtle point to note with respect to distributed deadlock detection is that delays in proI5agating local inforrnation rnight cause the deadlock detection algorithr11 to identify 'deadlocks' that do not really exist. Such situations, called phantoln **deadlocks**, lead to unnecessary aborts. For concreteness, we cliscuss the centralized algorithrn, although the hierarchical algorithrn suffers fr0111 the same problern.

Consider a rnodificatioll of the previous exarnple. As before, the two transactions wait for each other, generating the local \vaits-for graphs shown in Figure 22.5, and the local waits-for graphs are sent to the global deadlock-detection site. IIo\vever, $T2$ is now aborted for 1'easons other than deadlock. (For example, $T2$ rnay also be executing at a third site, 'where it reads an unexpected data value and decides to abort.) At this point, the local waits-for graphs have changed so that there is no cycle in the 'true' global \vaits-for graph. However, the constructed globaJ waits-for graph \vill contain a cycle, and $T1$ Inay well be picked as the victirn!

22.14 DISTRIBUTED RECOVERY

Recovery in a distributed DBMS is rnore cornplicated than In a centralized DBMS for the following reasons:

- New kinds of failure can arise: failure of COlnnlunication links and failure of a remote site at which a subtransaction is executing.
- Either all subtransactions of a given transaction Iuust ccnnlnit or none HUlst conlnlit, and this property IIlust be guaranteed despite any cOIllbination of site and link failures. T'his guarantee is achieved using a commit protocol.

As in a centralized DBMS, certain actions are carried out as part of norrnal execution to provide the necessary infonnation to recover fro111 failures. A log is rnaintained at each site, and in addition to the kinds of inforrnation rnaintained in a centralized DBMS, actions taken as part of the cOlInnit protocol are also logged. The Inost widely used conunit protocol is called *Two-Phase Cornmit (2PC)*. A variant called 21'C *with Presumed Abort*, which we discuss next, has been adopted as an industry standard.

In this section, we first describe the steps taken during nonnal execution, concentrating on the cOHnnit protocol, and tJlen discuss recovery fron1 failures.

22.14.1 Normal Execution and Commit I>rotocols

I)uring Donnal execution, each site rnaintains a log, and the actions of a subtransaction are logged at the site where it executes. The regular logging activity described in Chapter 18 is carried out and, in addition, a eornnlit protocol is followed to ensure that all subtra,nsa.ctions of a given transaction either cOIrnnit or H,bort uniforrnly. The transaction rnanager at the site where the transaction originated is called the coordinator for the transaction; transaction Inanagers at sites where its subtraJ1Sactiol1S execute are called subordinates (with respect to the coordinatioll of this transaction).

We no\v describe the **Two-Phase** Cornmit (2PC) protocol, in terms of the messages exchanged and the lc)g records written. When the user decides to cornrnit a transaction, the conunit cOlrunand is sent to the coordinator for the transaction. This initiates the 2PC protocol:

1. The coordinator sends a *prepare* rnessage to each subordinate.
2. When a subordinate receives a *prepare* message, it decides whether to abort or cornrnit its subtransaction. It force-writes an abort or prepare log record, and *then* sends a *no* or *yes* rnessage to the coordinator. Note that a prepare log record is not used in a centralized DBMS; it is unique to the distributed cornrnit protocol.
3. If the coordinator receives ***yes*** lnessages frOIII all subordinates, it force-writes a cornmit log record and then sends a *cornrnit* rnessage to all subordinates. If it receives even one *no* rnessage or receives no response fronl SOHle subordinate for a specified titne-out interval, it force-writes an abort log record, and then sends an *abort* Inessage to all subordinates.[1]
4. When a subordinate receives an *abort* Inessage, it force-writes an abort log record, sends an *ack* Inessage to the coordinator, and aborts the subtransaction. When a subordinate receives a *cornrnit* rnessage, it force-writes a cOl1nnit log record, sends an *ack* rnessage to the coordinator, and corrunits the subtransaction.
5. After the coordinator has received *ack* rnessages frorn all subordinates, it writes an end log record for the transaction.

1:'he narne *Two-Phase Commit* reflects the fact that two rounds of rnessages are exchanged: first a voting phase, then a tennination phase, both initiated by the coordinator. ffhe basic principle is that any of the transaction tnanagel'S involved (including the coordinator) can unilaterally a,bort a transaction, \vhereas therernust be unanirnity to conuuit a transaction, When a rnessage is serlt in 2PC, it signals a decision by the sender. To ensure that this decision survives a crash at the sender's site, the log record describing the decision is always forced to stable storage *before* the rnessage is sent.

;\ transaction is ofIicially cornrnitted at the tirne the coordiIlator's cOllnnit log record reaches stable storage. Subsequent failures cannot affect the outcorne of the transaction; it is irrevocaJ)ly corrunitted. Log records\vritten to record the connnit protocol actions contain the type of the record, the transaction id, and the identity of the coordinator. *i* coordinator's conunit or abort log record also contains the identities of the subordinates.

1As a,n optilnization, the coordinator need not send *abort* messages to subordinates who voted *no,*

22.14.2 Restart after a Failure

When a site COUles back up after a crash, we invoke a recovery process that reads the log and processes all transactions executing the conunit protocol at the tirne of the crash. The transaction rnanager at this site could have been the coordinator for SOUle of these transactions and a subordinate for others. We do the following in the recovery process:

- If we have a cornInit or abort log record for transaction T, its status is clear; we redo or undo T, respectively. If this site is the coordinator, which can be deterrnined froru the cOlTllnit or abort log record, we rnust periodically resend—because there rnay be other link or site failures in the system—a *commit* or *abort* rnessage to each subordinate until we receive an *ack*. After we have received *ack*s frorn all subordinates, we write an end log record for T.

- If we have a prepare log record for T but no conunit or abort log record, this site is a subordinate, and the coordinator can be detennined froIn the prepare record. We rllust repeatedly contact the coordinator site to determine the status of T. Once the coordinator responds with either cOlurnit or abort, we write a corresponding log record, redo or undo the transaction, and then write an end log record for T.

- If we have no prepare, cOllunit, or abort log record for transaction T, T certainly could not have voted to connuit before the crash; so we can unilaterally abort and undo T and write an end log record. In this case, we have no way to detennine whether the current site is the coordinator or a subordinate for T. flowever, if this site is the coordinator, it rnight have sent a *prepare* rnessage prior to the crash, and if so, other sites rnay have voted *yes*. If such a subordinate site contacts the recovery process at the current site, we now know that the current site is the coordinator for T, and given that there is no cOllnnit or abort log record, the response to the subordinate should be to abort T.

()bserve that, if the coordinator site for a transaction T fails, subordinates who voted *yes* cannot decide whether to conunit or abort T until the coordinator site recovers; we say that T is blocked. In principle, the active subordinate sites could cOl1nnunicate arnong thelIlselves, and if at least one of thelIl contains an abort or coinrnit log record for T, its status becornes globally known. 1"0 conununicate arnong thernselves, all subordinates nlust be told the identity of the other subordinates at the time they are sent the *prepare* message. llowever, 2PC is still vulnerable to coordinator failure durirlg recovery because even if all subordinates voted *yes*, the coordinator (who also has a vote!) may have decided to aJ)ort T, and this decision cannot be determined until the coordinator site recovers.

We covered how a site recovers fro111 a crash, but what should a site that is involved in the cOIIunit protocol do if a site that it is cornrl1unicating with fails? If the current site is the coordinator, it should shnply abort the transaction. If the current site is a subordinate, and it has not yet responded to the coordinator's *prepare* message, it can (and should) abort the transaction. If it is a subordinate and has voted *yes*, then it cannot unilaterally abort the transaction, and it cannot cOIUlnit either; it is blocked. It lnust periodically contact the coordinator until it receives a reply.

Failures of COffilnunication links are seen by active sites as failure of other sites that they are comnlunicating with, and therefore the solutions just outlined apply to this case as \-vell.

22.14.3 Two-Phase Commit Revisited

Now that we examined how a site recovers frolll a failure, and saw the interaction between the 2PC protocol and the recovery process, it is instructive to consider how 2PC can be refined further. In doing so, we arrive at a more efficient version of 2PC, but equally irnportant perhaps, we understand the role of the various steps of 2PC ruore clearly. Consider three basic observations:

1. l'he *ack* rnessages in 2PC are used to detennine when a coordinator (or the recovery process at a coordinator site following a crash) can 'forget' about a transaction *T*. lJntil the coordinator knows that all subordinates are aware of the cornrnit or abort decision for *T*, it IIlust keep inforrnation about *T* in the transaction table.

2. If the coordinator site fails after sending out *prepare* rnessages but before writing a cornrnit or abort log record, when it cornes back up, it has no inf'orruatiol1 abollt the transaction's connnit status prior to the crash. However, it is still free to abort the transaction unilaterally (beca,use it has not \vrittcn a conunit record, it can still cast a *no* vote itself). If another site inquires about the status of the transaction, the recovery process, as we have seen, responds \vith an *abort* rnessage. Therefore, in the absence of inforrnation, a transaction is *presumed to h..ave aborted.*

3. If a subtransaction does no updates, it has no changes to either redo or undo: in other words, its cornrnit or abort status is irrelevant.

The first two ol)servations suggest several refinements:

- When a coordinator aborts a transaction T, it can undo T and rerl10ve it fronl the transaction table irnrrlediately. After all\ rernoving T frorn the table results in a 'no inforrnatioI1' state with respect to T, and the default

response (to an enquiry about *T)* in this state, \vhich is *abort*, is the correct response for an aborted transaction.

- By the same token, if a subordinate receives an *abort* Inessagc, it need not send an *ack* lnessage. rrhe coordinator is not waiting to hear frorn subordinates after sending an *abor't* 111essage! If, for SOlne reason, a subordinate that receives a *prepare* message (and voted *yes)* does not receive an *abort* or *commit* Inessage for a specified tirne-out interval, it contacts the coordinator again. If the coordinator decided to abort, there Inay no longer be an entry in the transaction table for this transaction, but the subordinate receives the default *abort* nlessage, whicll is the correct response.

- Because the coordinator is not waiting to hear froul subordinates after deciding to abort a transaction, the names of subordinates need not be recorded in the abort log record for the coordinator.

- All abort log records (for the coordinator as well as subordinates) can simply be appended to the log tail, instead of doing a force-write. After all, if they are not written to stable storage before a crash, the default decision is to abort the transaction.

The third basic observation suggests SOlne additional refinements:

- If a subtransaction does no updates (which can be easily detected by keeping a count of update log records), the subordinate can respond to a *prepare* 111essage from the coordinator with a *reader* message, instead of *yes* or *no.* The subordinate writes no log records in this case.

- When a coordinator receives a *reader* lnessage, it treats the Inessage as a *yes* vote, but with the optiInization that it does not send any lnore messages to the subordinate, because the subordinate's cornlnit or abort status is irrelevant.

- If all subtransactions, including the sllbtransaction at the coordinator site, send a *reader* luessagc, we do not need the second phase of the conunit protocol. Indeed, we can sirnply rernove the transaction frolH the transaction table, \vithout \vriting any log records at any site for this transaction.

The T'wo-Phasc Cornrnit protocol with the refinernents discussed in this section is called Two-Phase Commit with Presurned Abort.

22.14.4 Three-Phase Commit

A cornlnitprotocol called Three-Phase Conlrnit (3PC) can avoid blocking even if the coordinator site fails during recovery. The basic idea is that, when

the coordinator sends out *prepare* rnessages and receives *yes* votes £1'0111 all subordinates, it sends all sites a *precommit* message, rather than a *cornrnit* rnessage. When a sufficient l1Ulllber.. ..more than the lIlaxinlulll nUlnber of failures that nlust be handled·····o*acks* have been received, the coordinator force-writes a *cornrn·it* log record and sends a *cornmit* lnessage to all subordinates. In 3PC, the coordinator effectively postpones the decision to cornrnit until it is sure that enough sites know about the decision to corn111it; if the coordinator subsequently fails, these sites can C0l11111Unicate with each other and detect that the transaction rnust be corllrnitted-conversely, aborted, if none of thern has received a *precomrnit* rnessage-'-without waiting for the coordinator to recover.

rrhe 3PC protocol ilnposes a significant additional cost during normal execution and requires that COlnrIlunication link failures do not lead to a network partition (wherein sorne sites cannot reach some other sites through any path) to ensure freedo111 fronl blocking. For these reasons, it is not used in practice.

22.15 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Discuss the different rnotivations behind parallel and distributed databases. (Section 22.1)
- Describe the three lllain architectures for parallel DBMSs. Explain why the *shared-memory* and *shared-disk* approaches suffer frOlll *interference*. What can you say about the *speed-up* and *scale-up* of the *shared-nothing* architecture? (Section 22.2)
- Describe and differentiate *pipelined parallelism* and *data-partitioned parallelism*. (Section 22.3)
- Discuss the following techniques for partitioning data: *round-Tobin, hash,* and *range*. (Section 22.3.1)
- Explain how existing code can be parallelized by introducing *split* and *merge* operators. (Section 22.3.2)
- Discuss how each of the following operators can be parallized using data partitioning: *scanning, sorting, join*. Cornparc the use of sorting versus hashing for partitioning. (Section 22.4)
- What do we need to consider in optilllizing queries for parallel execution? Discuss interoperation parallelislll, left-deep trees versus bushy trees, and cost estimation. (Section 22.5)

- Define the tenns *distributed data independence* and *distributed transaction atomicity.* Are these concepts sllpported in current eornrnercial systerns? \Vhy not? What is the difference between *homogeneous* and *heterogeneous* distributed databases? (Section 22.6)

- Describe the three lllain architectures for distributed DBMSs. (Section 22.7)

  1\ relation can be distributed by *fragmenting* it or *replicating* it across several sites. Explain these concepts and ho\v they differ. Also, distinguish between *horizontal* and *vertical* fragrnentation. (Section 22.8)

- If a relation is fraglnented and replicated, each partition needs a globally unique nalne called the *Tclat'ion narnc.* Explain how such global naInes are created and the Inotivation behind the described approach to narning. (Section 22.9.1)

- Explain how rnetadata about such distributed data is rnaintained in a *distr'ibuted catalog.* (Section 22.9.2)

- Describe a nauling scherne that supports distributed data independence. (Section 22.9.3)

- When processing queries in a distributed DBMS, the location of partitions of the relation needs to be taken into account. Discuss the alternatives when joining two two relations that reside on different sites. In particular, explain and describe the rnotivation behind the *Sernijoin* and *Bloornjoin* techniques. (Section 22.10.2)

- What issues rnust be considered in optirnizing queries over distributed data, in addition to where the data is located? (Section 22.10.3)

- What is the difference bet\veen *synchronous asynchronous* replication? Why has asynchronous replication gained in popularity? (Section 22.11)

- Describe the *'ooting* and *Tead-a'ny write-all* approaches to synchronous replication. (Section 22.11.1)

- Surnruarize the *peer-to-peer* and *primary site* approaches to asynchronolls replication. (Section 22.11.2)

- In prirnary site replication, changes to the prirnary copy Inust be propagated to secondary copies. What is done in the *Capture* and *Apply* steps? Describe *log-based* and *procedural* approaches to Capture and cornpare theln. What are the variations in scheduling the Apply step? Illustrate the use of asynchronolls replication in a data warehouse. (Section 22.11.2)

- What is a *subtransaction?* (Section 22.12)

- What are the choices for rnanaglng locks in a distributed DBMS? (Section 22.13)
- Discuss deadlock detection in a distributed database. Contrast the *centralized*, *hierarchical*, and *time-out* approaches. (Section 22.13.1)
- Why is recovery in a distributed DBMS rIlore cornplicated tha.n III a centralized systern? (Section 22.14)
- What is a *commit protocol* and \vhy is it required in a distributed database? Describe and compare *Two-Phase* and *Three-Phase Commit*. What is *blocking*, and how does the Three-Phase protocol prevent it? Why is it nonetheless not used in practice? (Section 22.14)

EXERCISES

Exercise 22.1 Give brief answers to the following questions:

1. What are the siruilarities and differences between parallel and distributed database rnan-· agement systerns?
2. Would you expect to see a parallel database built using a wide-area network? \Vould you expect to see a distributed database built using a wide-area network? Explain.
3. Define the terms *scale-up* and *speed-up*.
4. Why is a shared-nothing architecture attractive for parallel database systerns?
5. The idea of building specialized hardware to run parallel database applications received considerable a[illegible]tion but has fallen out of favor. Cornrnent on this trend.
6. What are th,[illegible]antages of a distributed D131\1[8 over a centralized DBMS?
7. Briefly descr[illegible]and cornpare the Client-Server and Collaborating Servers architectures.
8. In the Colla[illegible]ting Servers architecture, \vhen a transaction is subrnitted to the DBMS, briefly des[illegible]how its activities at various sites are coordinated. In particular, describe the role o[illegible].saction managers at the different sites, the concept of *subtransactions*, and the,[illegible]cept of *distributed transaction atomicity*.

Exercise 22.2 Give brief a,nswers to the follmving questions:

1. I)efine the tenus *fragmentation* and *replicatiOn* in tenns of where data is stored.
2. Wha' is the difference between *synchronous* and *asynchronous* replication?
3. Define the tern1 *distributed data independence*. What does this Inean'with respect to quer:ying and updating data in the presence of data fragrnentation and replication?
4. Consider the *voting* and *read-any write-all* techniques for irnplementing synchronous replication. What are their respective pros and cons?
5. Give an overview of how asynchronous replication can be implemented. In particular, explain the terms *Capture* and *Apply*.
6. What is the difference between log-based and procedural irnplernentatiOlls of capture?
7. Why is giving database objects unique names rnore cennplicated in a distributed DBMS?

8. Describe a catalog organization that pennits any replica (of an entire relation or a frag-rnent) to be given a unique nam,e and provides the nanling infrastructure required for ensuring distributed data independence.
9. If infonuation from renlote catalogs is cached at other sites, what happens if the cached infoflllation becOInes outdated? How can this condition be detected and resolved?

Exercise 22.3 Consider a parallel DBMS in \vhich each relation is stored by horizontally partitioning its tuples across all disks:

Ernployees(eid: integer, *did:* integer, *sal:* real)
Departments(*did:* integer, *mgrid:* integer, *budget:* integer)

The *mgrid* field of DepartInents is the *e'id* of the manager. Each relation contains 20-byte tuples, and the *sal* and *budget* fields both contain unifonnly distributed values in the range 0 to 1 rnillion. The Enlployees relation contains 100,000 pages, the Departrnents relation contains 5,000 pages, and each processor has 100 buffer pages of 4,000 bytes each. The cost of one page I/O is t_d, and the cost of shipping one page is t_s; tuples are shipped in units of one page by waiting for a page to be filled before sending a rnessage frmn processor i to processor j. 'There are no indexes, and all joins that are local to a processor are carried out using a sort-rnerge join. Assurne that the relations are initially partitioned using a round-robin algorithlll and that there are 10 processors.

For each of the following queries, describe the evaluation plan briefly and give its cost in tenns of t_d and t_s. You should cornpute the total cost across all sites as well as the 'elapsed time' cost (i.e., if several operations are carried out concurrently, the tirne taken is the rnaxilnurn over these operations).

1. Find the highest paid ernployee.
2. Find the highest paid employee in the departrnent with *did 55.*
3. Find the highest paid ernployee over all departnHmts with *budget* less than 100,000.
4. Find the highest paid enlployee over all departlnents with *budget* less than 300,000.
5. Find the a;verage salary over all departments with *budget* less than 300,000.
6. Find the salaries of all rnanagers.
7. Find the salaries of all rnanagers who manage a departrnent with a budget less than 300,000 and eaTll rnore than 100,000.
8. Print the *eids* of all elnployees, ordered by increasing salaries. Each processor is connected to a separate printer, and the answer can appear as several sorted lists, each printer] by a different processor, as long as we can ol)tain a fully sorted list by concatenating the printed lists (in sorne order).

Exercise 22.4 Consider the saIne scenario as in Exercise 22.3, except that the relations are originally partitioned using range partitionirlg on the *sal* and *budget* fields.

Exercise 22.5 Repeat Exercises 22.3 and 22.4 with (i) 1 processor, and (ii) lelO processors.

Exercise 22.6 COllsicler the Ernployees and Departments relations described in Exercise 22.3. They are now stored in a distributed DBMS with all of Employees stored at Naples and all of Departlnents stored at Berlin. There arc no indexes on these relations. The cost of various operations is as describecl in Exercise 22.:3. Consider the query:

```
SELECT *
FROM   EInployees E, Dcpartrncnts D
WHERE  E.eid = I).Ingrid
```

'The query is posed at Delhi, and you are told that only 1 percent of ernployees are IIlanagers. Find the cost of answering this query using each of the following plans:

1. Ship Departrnents to Naples, cornpute the query at Naples, then ship the result to Delhi.
2. Ship Ernployees to Berlin, cornpute the query at Berlin, then ship the result to Delhi.
3. COInpute the query at Delhi by shipping both relations to Delhi.
4. COlnpute the query at Naples using BloOlnjoin; then ship the result to Delhi.
5. Compute the query at Berlin using Bloornjoin; then ship the result to Delhi.
6. Cornpute the query at Naples using Sernijoin; then ship the result to Delhi.
7. COInpute the query at Berlin using Sernijoin; then ship the result to Delhi.

Exercise 22.7 Consider your answers in Exercise 22.6. Which plan ll1inin1izes shipping costs? Is it necessarily the cheapest plan? Which do you expect to be the cheapest?

Exercise 22.8 Consider the Ernployees and Departments relations described in Exercise 22.3. They are now stored in a distributed DBMS with 10 sites. The DepartInents tuples are horizontally partitioned across the 10 sites by *did,* with the same nUInber of tuples assigned to each site and no particular order to how tuples are assigned to sites. The Employees tuples are sirnilarly partitioned, by *sal* ranges, with $sal \leq 100{,}000$ assigned to the first site, $100{,}000 < sal \leq 200{,}000$ assigned to the second site, and so Oll. In addition, the partition $sal \leq 100{,}000$ is frequently accessed and infrequently updated, and it is therefore replicated at every site. No other EU1ployees partition is replicated.

1. Describe the best plan (unless a plan is specified) and give its cost:
   (a) Cornpute the natural join of Enlployees and Departlnents by shipping all fragrnents of the slImller relation to every site containing tuples of the larger relation.
   (b) Find the highest paid ernployee.
   (c) Find the highest paid clnployee with salary less than 100,000.
   (d) Find the highest paid ernployee with salary between 400,000 and 500,000.
   (e) 'Find the highest paid clnployee with salary between 450,000 and 550,000.
   (f) Find the highest paid rnanager for those departnwnts stored at the query site.
   (g) Find the highest pajd lnanager.
2. ASSU111ing the sarne data distribution, describe the sites visited and the locks obtained for the foll()\ving update transactions, assuming that *synchronous* replication is used for the replication of Employees tuples with $sal \leq 100,(){}{}$:
   (a) Give employees with salary less than 100,000 a 10 percent raise, with a Inaxirnurn salary of 100,000 (i.e., the raise cannot increase the salary to rnore than 100,(00).
   (b) Give all ernployees a 10 percent raise. The conditions of the original partitioning of Elnployees IIlust still be satisfied after the update.
3. AssuIning the saIne data distribution, describe the sites visited and the locks obtained for the following update transactions, assuming that *asynchronous* replication is used for the replication of Ernployees tuples with $sal \leq 100{,}000$.

For all employees with salary less than 100,000 give them a 10 percent raise, with a maximum salary of 100,000.

Give all employees a 10 percent raise. After the update is completed, the conditions of the original partitioning of Eluployecs rnust still be satisfied.

Exercise 22.9 Consider the EInployees and Departments tahles from Exercise 22.:3. You are a DBA and you need to decide how to distribute these two tables across two sites, IVlanila and Nairobi. Your DBMS supports only unclustered 13+ tree indexes. You have a choice between synchronous and asynchronous replication. For each of the following scenarios, describe how you would distribute thenl and what indexes you would build at each site. If you feel that you have insufficient information to make a decision, explain briefly.

1. Half the departInents are located in Manila and the other half are in Nairobi. Departrnent information, including that for ernployees in the depart1nent, is changed only at the site where the departrnent is located, but such changes are quite frequent. (Although the location of a depart1nent is not included in the Departrnents schclna, this inforrnation can be obtained frorn another table.)
2. Half the departrnents are located in Manila and the other half are in Nairobi. Departrnent information, including that for errlployees in the departrnent, is changed only at the site where the departrnent is located, but such changes are infrequent. F'inding the average salary for each departrnent is a frequently asked query.
3. Half the departlnents are located in Ivlanila and the other half are in Nairobi. Ernployees tuples are frequently changed (only) at the site where the corresponding departrrlent is located, but the Depart1nents relation is aJulOst never changed. Finding a given ernployee's rnanager is a frequently asked query.
4. Half the e1nployees work in Manila and the other half work in Nairobi. E1nployees tuples are frequently changed (only) at the site where they work.

Exercise 22.10 Suppose that the Ernployees relation is stored in I\1adison and the tuples with $sal \leq 1.00{,}000$ are replicated at New York. Consider the following three options for lock rnanagernent: all locks managed at a *single site*, say, 1VIilwaukee; *primary copy* with l'vladison being the primary for Employees; and *fully distributed*. For each of the lock rnanagernent options, explain what locks are set (and at which site) for the following queries. Also state frorn which site the page is reac1.

1. A query at Austin wants to read a page of Erllployees tuples \vith $sal \leq 50{,}0OO$.
2. A query at Madison wants to read a page of Employees tuples with $sal \leq 50{,}000$.
3. A query at New York wants to read a page of Employees tuples "vith $sal \leq 50{,}000$.

Exercise 22.11 Briefly answer the following questions:

1. Compare the relative merits of centralized and hierarchical deadlock detection in a distril)lIted DBMS.
2. What is a *phantom deadlock*? Give an example.
3. Give an example of a distributed DBMS \vith three sites such that no two local waits-for graphs reveal a deadlock, yet there is a global deadlock.
4. Consider the following rnoclification to a local waits-for graph: Add a new node T_{ext}, and for every transaction T_i that is waiting for a lock at another site, add the edge $T_i \quad T_{ext}$. A.lso add an edge $T_{ext} \rightarrow T_i$ if a transaction executing at another site is waiting for T_i to release a lock at this site.

If there is a cycle in the modified local waits-for graph that does not involve T_{ext}, what can you conclude? If every cycle involves T_{ext}, what can you conclude?

Suppose that every site is assigned a unique integer Whenever the local waits-for graph suggests that there Blight be a global deadlock, send the local waits-for graph to the site with the next higher site-id. At that site, combine the received graph with the local waits-for grap,h. If this cornbined graph does not indicate a deadlock, ship it on to the next site, and so on, until either a cleadlock is detected or we are back at the site that originated this round of deadlock detection. Is this scheme guaranteed to find a global deadlock if one exists?

Exercise 22.12 Tirnestarnp-based concurrency control schernes can be used in a distributed DBMS, but we rllust be able to generate globally unique, rllonotonicaJly increasing timestamps without a bias in favor of anyone site. One approach is to assign timestamps at a single site. Another is to use the local clock tiTne and to append the site-iei. A third scherne is to use a counter at each site. COIllpare these three approaches.

Exercise 22.13 Consieler the rIlultiple-granlllarity locking protocol described in Chapter 18. In a distributed DBMS, the site containing the root object in the hierarchy can becmne a bottleneck. You hire a database consultant who tells you to rnodify your protocol to allow only intention locks on the root and irnplicitly grant all possible intention locks to every transaction.

1. Explain why this rnodification \vorks correctly, in that transactions continue to be able to set locks on desired parts of the hierarchy.
2. Explain how it reduces the demand on the root.
3. Why is this idea not included as part of the standard rllultiple-granularity locking protocol for a centralized DBMS?

Exercise 22.14 Briefly answer the following questions:

1. Explain the need for a cornmit protocol in a distributed DBMS.
2. Describe 2PC. Be sure to explain the need for force-writes.
3. Why are *ack* messages required in 2PC?
4. What are the differences between 2PC and 2PC with PresulTled Abort?
5. Give an exarnple execution sequence such that 2PC and 2PC 'with Presurned Abort: generate an identical sequence of actions.
6. Give an exarIlple execution sequence such that 2PC and 2PC with Presumed Abort generate different sequences of actions.
7. \Vhat is the intuition behind 3PC? What are its pros and cons relative to 2PC?
8. Suppose that a site gets no response frorn another site for a long time. Can the first site tell whether the connecting link has failed or the other site has failed? How is such a failure handled?
9. Suppose that the coordinator incluc1es a list of aU subordinates in the *prepare* message. If the coordinator fails after sending out either an *abort* or *commit* message, call you suggest a way for active sites to terrninate this tra.nsaction without wajting for the coordinator to recover? Assume that some but not all of the *abort* or *commit* messages frOln the cocn'clinator are lost.

III Suppose that 2PC with PresuIl1ed Abort is used as the cOInmit protocol. Explain how the systmll recovers froIn failure and deals with a particular transaction T in each of the following cases:

(a) A subordinate site for T fails before receiving a *prepare* rnessage.

(b) A subordinate site for T fails after receiving a *prepare* rnessage but before rnaking a decision.

(c) A subordinate site for T fails after receiving a *prepare* lnessage and force-writing an abort log record but before responding to the *prepare* message.

(d) A subordinate site for T fails after receiving a *prepare* message and force-writing a prepare log record but before responding to the *prepare* lnessage.

(e) A subordinate site for T fails after receiving a *prepare* rnessage, force-writing an abort log record, and sending a *no* vote.

(f) The coordinator site for T fails before sending a *prepare* lnessage.

(g) The coordinator site for T fails after sending a *prepare* lllCssage but before collecting all votes.

(h) The coordinator site for T fails after writing an *abort* log record but before sending any further rnessages to its subordinates.

(i) The coordinator site for T fails after writing a *comrnit* log record but before sending any further rnessages to its subordinates.

(j) The coordinator site for T fails after writing an *end* log record. Is it possible for the recovery process to receive an inquiry about the status of T frolll a subordinate?

Exercise 22.15 Consider a heterogeneous distributed DBMS.

1. Define the terms *multidatabase system* and *gateway.*
2. Describe how queries that span multiple sites are executed in a rnultidatabase systern. Explain the role of the gateway with respect to catalog interfaces, query optirnizatiOll, and query execution.
3. Describe how transactions that update data at rnultiple sites are executed in a lllulti-database systern. Explain the role of the gateway with respect to lock rnanagernent, distributed deadlock detection, Two-Phase COllnnit, and recovery.
4. SChell1aS at different sites in a llnI1tidatabase systern are probably designed independently. T'his situation can lead to *semantic heterogeneity;* that is, units of rneasure rnay differ across sites (e.g., inches versus centirneters), relatiolls containing essentially the same kind of infonnation (e.g., employee salaries and ages) rnay have slightly different schernas, and so on. What ilnpact does this heterogeneity have on the end user? In particular, COllunent on the concept of distributed data independence in such a systcrIl.

BIBLIOGRAPHIC NOTES

\Vork on parallel algorithrns for sorting and various relational operations is discussed in the bibliographies for Chapters 13 and 14. Our discussion of parallel joins follows [220], and our discussion of parallel sorting follows [223]. DeWitt and Gray make the case that for future high performance database systeuls, parallelisnl will be the key [221]. Scheduling ill parallel database systenls is discusserl in [522]. [496] contains a good collection of papers on query processing in parallel database systems.

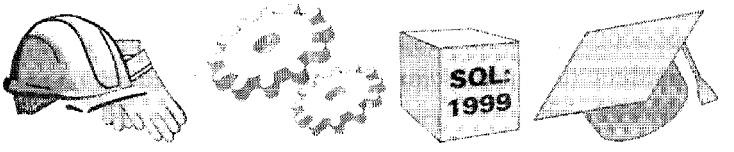
Textbook discussions of distributed databases include [78, 144, 580]. Good survey articles include [85], which focuses on concurrency control; [637], which is about distributed databases in general; and [785], which concentrates on distributed query processing. Two InajaI' projects in the area were 8DD-1 [636] and R* [777]. Fragmentation in distributed databases is considered in [157, 207]. Replication is considered in [11, 14, 137,239,238, 388, 385, 335, 552, 600]. For good overviews of current trends in asynchronous replicatioll, see [234, 709, 772]. Papers on view maintenance mentioned in the bibliographic notes of Chapter 21 are also relevant in this context. Olston considers techniques for trading of performance versus precision in a replicated environment [571, 572, 573].

Query processing in the 8DD-1 distributed database is described in [88]. One of the notable aspects of 8DD-1 query processing was the extensive use of 8emijoins. Theoretical studies of Semijoins are presented in [83, 86, 414]. Query processing in R* is described in [667]. The R* query optimizer is validated in [500]; much of our discussion of distributed query processing is drawn from the results reported in this paper. Query processing in Distributed Ingres is described in [247]. Optimization of queries for parallel execution is discussed in [297,323,383]. Franklin, Jonsson, and Kossman discuss the trade-offs between *query shipping,* the more traditional approach in relational databases, and *data shipping,* which consists of shipping data to the client for processing and is widely used in object-oriented systerns [284]. A good recent survey of distributed query processing techniques can be found in [450].

Concurrency control in the 8DD-1 distributed database is described in [91]. Transaction management in R* is described in [547]. Concurrency control in Distributed Ingres is described in [714]. [740] provides an introduction to distributed transaction rrlanagement and various notions of distributed data independence. Optimizations for read-only transactions are discussed in [306]. Multiversion concurrency control algorithrlls based on timestamps were proposed in [620]. Timestamp-based concurrency control is discussed in [84, 356]. Concurrency control algorithms based on voting are discussed in [303, 318, 408, 452, 732]. The rotating prirnary copy scheme is described in [538]. Optimistic concurrency control in distributed databases is discussed in [660], and adaptive concurrency control is discussed in [488].

Two-Phase Commit was introduced in [466, 331]. 2PC with Presumed Abort is described in [546], along with an alternative called *2PC with Presum.ed Cornmit.* A variation of Presumed Cornrrlit is proposed in [465]. Three-Phase COlnrrlit is described in [692]. The deadlock detection algorithnls in R* are described in [567]. Many papers discuss deadlocks, for exaInple, [156, 243, 526, 632]. [441] is a survey of several algoritluns in this area. Distributed clock synchronization is discussed by [464]. [333] argues that distributed data independence is not always a good idea, clue to processing and adlninistrative overheads. The ARIES algorithrl1 is applicable for distributed recovery, but the details of how rnessages should be handled are not discussecl in [544]. The approach taken to recovery in SDD-1 is described in [43]. [114] also addresses distributed recovery. [444] is a survey article that discusses concurrency control and recovery in distributed systerIls. [95) contains several articles on these topics.

IVlultidatabase systems are discussed in [10, 113, 230, 231, 242,476, 485, 519, 520, 599, 641, 765, 797]; sec [112, 486, 684] for surveys.



23

OBJECT-DATABASE SYSTEMS

- What are object-database systerlls and what new features do they support?
- What kinds of applications do they benefit?
- \Vhat kinds of data types can users de.fine?
- What are abstract data types and their benefits?
- What is type inheritance and why is it useful?
- What is the irnpact of introducing object ids in a database?
- How can we utilize the new features in database design?
- What are the new implelncntation challenges?
- \Vhat difFerentiates object-relational and object-oriented DBMSs?
- Key concepts: user-defined data types, structured types, collection types; data abstraction, rnethocls, encapsulation; inheritance, early and late binding of rnethods, collection hierarchies; object identity, reference types, shallow and deep equality

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University of California-Berkeley

You know Iny Inethods, Watson. Apply theIn.

Arthur Conan Doyle, *The Memoirs of Sherlock Holmes*

Relational database systeros support a sruaU, fixed collection of data types (e.g., integers, dates, strings),\vhich has proven adequate for traditional application domains such as adruinistrative data processing. In many application dornains, however, rnuch 1nore complex kinds of data Blust be handled. Typically this cornplex data has been stored in OS file systems or specialized data structures, rather than in a DBMS. Examples of dornains with cOJ.uplex data include cornputer-aided design and rnodeling (CA.D/CAM), multimedia repositories, and docurnent Hl8.Jlagernent.

As the arnount of data grows, the luany features offered by a DBIvISfor exarIlpIe, reduced application developnlent time, concurrency control and recovery, indexing support, and query capabilities·····becoruincreasingly attractive and, ultimately, necessary. To support such applications, a DBMS HUlst support cornplex data types. ()bject-oriented concepts strongly influenced efforts to enhance database support for cornplex data and led to the developrnent of object-database systelus, \vhich we discuss in this chapter.

Object-database systerlls have developed along two distinct paths:

- Object-Oriented Database Systems: Object-oriented database systerns are proposed as an alternative to relational systerlls and are ainled at application dornains where cODlplex objects playa centra,} role. The approach is heavily influenced by object-oriented prograrllrlling languages and can be understood as an atternpt to add DBMS functionality to a prograunning language environrnent. The ()bject Database :M:anagenlcnt Group (()DMG) has developed a standard Object Data Model (ODM) and Object Query Language (OQL), which are the equivalent of the SQL standard for relational database systerns.
- Object-Relational])atabase Systenls: ()bject-relational database s.systerns ca,n be thought of as an atternpt to extend relational database systerns with the functionality necessary to support a broader class of applications and, in many ways, provide a bridge between the relational and object-oriented paTadiguls. The SQL:1999 standard extends SQL to incorporate support for the ol)ject-relationaJ rnode1 of data.

We use acronyms for relational, object-oriented, and object-relational database rnanagernent systems (RDBMS, OODBMS, ORJDBMS). In this chapter, we focus 011 ORDBMSs and ernphasize how they can be viewed as a developrnent of HJ)BMSs, rather than as an entirely different paradigrn, as exernplified l)y the evolution of SQL:1999.

We concentrate on developing the fUlldarnental concepts rather than presenting SQL:1999; some of the features we discuss are not inc.luded in SQL:1999.

Nonetheless, we have chosen to emphasize concepts relevant to SQL:1999 and its likely future extensions. We also try to be consistent with SQL:1999 for notation, although we occasionally diverge slightly for clarity. It is hnportant to recognize that the rnain concepts discussed are COlllllTIOn to both ORDBMSs and ()()DBNISs; we discuss how they are supported in the ODLjOQL standard proposed for OODBMSs in Section 23.9.

RDB1\JIS vendors, including IBM, Inforrnix, and Oracle, are adding ORDBMS functionality (to varying degrees) in their products, and it is inlportant to recognize how the existing body of knowledge about the design and inlplernnentation of relational databases can be leveraged to deal with the ORDBMS extensions. It is also ilnportant to understand the challenges and opportunities these extensions present to database users, designers, and irnplernentors.

In this chapter, Sections 23.1 through 23.6 introduce object-oriented concepts. The concepts discussed in these sections are COlunlon to both OODBMSs and ORDBJVISs. We begin by presenting an example in Section 23.1 that illustrates why extensions to the relational rnodel are needed to cope with some new application dornains. 'This is used as a running exarnple throughout the chapter. We discuss the use of type constructors to support user-defined structured data types in Section 23.2. We consider what operations are supported on these new types of data in Section 23.3. Next, we discuss data encapsulation and abstract data types in Section 23.4. We cover inheritance and related issues, such as rnethod binding and collection hierarchies, in Section 23.5. We then consider objects and object identity in Section 23.6.

We consider how to take advantage of the new object-oriented concepts to do OI{DBMS database design in Section 23.7. In Section 23.8, we discuss SOHle of the new irnplernentation challenges posed by object-relational systerns. We discuss ()I)L and OQL, the standards for OODBMSs, in Section 23.9, and then present a brief cornparison of ()R,DBMSs and OODBMSs in Section 23.10.

23.1 MOTIVATING EXAMPLE

As a specific exarnple of the need for object-relational systcrlls, we focus on a new business data processing probler.n that is both harder and (in our view) rnorc entertaining than the dollars and cents bookkeeping of previous decades. Today, cornpanies in industries such as entertainruent are in the business of selling *bits;* their basic corporate assets are not tangible products, but rather software artifacts such as video (l,nd audio.

We consider the fictional Dinky Entertaiurnent Corupa,ny, a large IIollywood conglornerate whose main assets are a collection of cartoon characters, espe-

cially the cuddly and internationally beloved IIerbert the \VarIll. Dinky has several IIerbert the \Vornlfihns, rnany of which are shown in theaters around the world at any given time. Dinky also rnakes a good deal of rnoney licensing Herbert's irnage, voice, and video footage for various purposes: action figures, video gaInes, product endOrSelllents, and so on. I)inky's database is used to lnanage the sales and leasing records for the various IIerbert-related products, as well as the video and audio data that rnake up IIerbert's lllany filIns.

23.1.1 New **Data** Types

The basic problern confronting Dinky's database designers is that they need support for considerably richer data types than is available in a relational DBMS:

- User-defined data types: Dinky's assets include Herbert's iIllage, voice, and video footage, and these rnust be stored in the database. To handle these new types, we need to be able to represent richer structure. (See Section 23.2.) Further, we need special functions to rnanipulate these objects. For example, we may want to write functions that produce a cOlnpressed version of an irnage or a lower-resolution image. By hiding the details of the data structure through the functions that capture the behavior, we achieve *data abstract'ion,* leading to cleaner code design. (See Section 23.4.)
- Inheritance: As the nurnber of data types grows, it is irnportant to take advantage of the cornrnonality between different types. :For exarnple, both cOInpressed irnages and lower-resolution irnages are, at SOlne level, just ilnages. It is therefore desirable to *inherit* some features of iluage objects while defining (and later Inanipulating) cOInpressed irnage objects and lower-resolution irnage objects. (See Section 23.5.)
- Object Identity: Given that seHne of the new data types contain very large instances (e.g., videos), it is iInportant not to store copies of objects; instead, we IIIUSt store *references*, or *pointers*, to such objects. In turn, this underscores the need for giving objects a unique *object identity*, which can be used to refer or 'point' to theln frorn elsewhere in the data. (See Section 23.6.)

Flow lnight we address these issues in an RDBMS? We could store ilnages, videos, and so on as BLC)Bs in current relational systems. A binary large object (BLOB) is just a long stream of bytes, and the DBMS's support consists of storing and retrieving BLC)Bs in such a rnanner that a user does not have to worry about the size of the BLC)B; a 13LC}B can span several pages, unlike a traditional attribute. All further processing of the BLC)B has to be done by the user's application prograrnl, in the host language in \vhich the

The SQL/MM Standard: SQL/MM is an eillerging standard that builds upon SQL:1999's new data types to define extensions of SQL:1999 that facilitate handling of coruplex multimedia data types. SQL/MM is a rnultipart standard. Part 1, SQL/MM Framework, identifies the SQL:1999 concepts that are the foundation for SQL/MM extensions. Each of the relnaining parts addresses a specific type of ccnnplex data: Full Text, Spatial, Still Image, and Data Mining. SQL/MM anticipates that these new coruplex types can be used in colurnns of tables as field values.

Large Objects: SQL:1999 includes a new data type called LARGE OBJECT or LOB, with two variaJ1ts called BLOB (binary large object) and CLOB (character large object). This standardizes the large object support found in lnany current relational DBMSs. LOBs cannot be included in priruary keys, GROUP BY, or ORDER BY clauses. rrhey can be cornpared llsing equa.lity, inequality, and substring operations. A LOB has a locator that is essentially a unique id and allows LOBs to be rnanipulated without extensIve copYIng.
L()Bs are typically stored separately froIn the data records in whose fields they appear. IBM DB2, InforInix, Microsoft SQL Server, Oracle 8, and Sybase ASE all support LOBs.

SQL code is ernbedded. This solution is not efficient because we are forced to retrieve all BLOBs in a collection even if rnost of the111 could be filtered out of the answer by applying user-defined functions (within the DBMS). It is not satisfactory frorn a data consistency standpoint either, because the selnantics of the data now depends heavily on the host la,nguage application code and cannot be enforced by the DBMS.

As for structured types and inheritance, there is simply no support in the relational Dlodel. We are forced to map data 'with such cOlnplex structure into a collection of flat tables. (We saw examples of such rnappings when we discussed the translation frorH ER diagrarns with inheritance to relations in Chapter 2.)

This application clearly requires features not available in the relational Inode1. As an illustration of these features, Figure 2:3.1 presents SQL:1999 :DDL staternents for a l)Ortion of Dinky's ()HJ)I31VIS schema used in subsequent examples. Although the 1)})L is very sirnilar to that of a traditional relational systeru, some irnportant distinctions highlight the new data rnodeling capabilities of an ORDBMS. A quick glance at the l)DL staternents is sufficient for now; we study them in detail in the next section, after presenting S0111e of the basic

concepts that our sanlple application suggests are needed in a next-generation DBMS.

1. CREATE TABLE Frames
   (*frameno* integer, *image* jpeg_image, *category* integer);
2. CREATE TABLE Categories
   (*cid* integer, *name* text, *lease_price* float, *comments* text);
3. CREATE TYPE theater_t AS
   ROW(*tno* integer, *name* text, *address* text, *phone* text)
   REF IS SYSTEM GENERATED;
4. CREATE TABLE Theaters OF theater_t REF is tid SYSTEM GENERATED;
5. CREATE TABLE Nowshowing
   (*film* integer, *theater* REF(theater_t) SCOPE rrheaters, *start* date,
   end date);
6. CREATE TABLE FilIns
   (filrnno integer, *title* text, *stars* VARCHAR(25) ARRAY [10]),
   director text, *budget* float);
7. CREATE TABLE Countries
   (*name* text, *boundary* polygon, *population* integer, *language* text);

Figure 23.1 SQL:1999 DDL Staternents for Dinky Schema

23.1.2 Manipulating the New Data

Thus far, we described the new kinds of data that rnust be stored in the Dinky database. We have not yet said anything about how to *use* these new types in queries, so let us study two queries that I)inky's database needs to support. The syntax of the queries is not critical; it is sufficient to understand what they express. We return to the specifics of the queries' syntax later.

()ur first challenge COIn8S frorn the Clog breakfast cereal cornpany. Clog produces a cereal called I)elirios and it wants to lease an irnage of Herbert the Worm in front of a sunrise to incorporate in the I)elirios box design. A query to present a collection of possible irnages and their lease prices can be expressed in SQL-like syntax as in Figure 2:3.2. I)inky has a nUJn1>er of methods written in an irnperative language like Java and registered with the database systern. These methods can be used in queries in the sallIe way as built-ill methods, such as =. , —, <, >, are used in a relational language like SQL. The *thumbnail* IJlethod in the Select clause produces a srnaU version of its full-size input: image. The *is_sunrise* rnethod is a boolean function that analyzes an irnage and returns *true* if the image contains a sun.rise; the *is_herbert* Inethod returns *true* if the image contains a picture Of l1erbert. rrhe query produces the frame

code number, irnage thurnbnail, and price for all frames that contain Herbert and a sunrise.

```
SELECT  F.fnuneno, thulnbnail(F.irnage), C.lease_price
FROM    Fralnes F, Categories C
WHERE   F.category = C.cid AND is.Bllnrise(F.irnage) AND isJlerbert(F.inlage)
```

Figure 23.2 Extended SQL to Find Pictures of Herbert at Sunrise

The second challenge carnes froIn Dinky's executives. They know that Delirios is exceedingly popular in the tiny country of A.ndorra, so they want to lIlake sure that a number of Herbert filIns are playing at theaters near Andorra when the cereal hits the shelves. To check on the current state of affairs, the executives want to find the 11aInes of all theaters showing Herbert fihns within 100 kilorneters of Andorra. Figure 23.3 shows this query in an SQL-like syntax.

```
SELECT  N.theater···-->na1ne, N.theater-> address, F.title
FROM    Nowshowing N, Filrns F, Countries C
WHERE   N.film = F.filrnno AND
        overlaps(C.bollndary, radius(N.theater-> address, 100)) AND
        C.name = 'Andorra' AND 'Herbert the Worm' = F.stars[l]
```

Figure 23.3 Extended SQL to Find Herbert Films Playing near Andorra

The *theater* attribute of the Nowshowing table is a reference to an object in another table, which has attributes *narne, address*, and *location.* This object referencing allows for the notation *N. theater-> narne* and *N. theater···>address,* each of which refers to attributes of the theater_t object referenced in the Nowshowing row *N.* The *stars* attribute of the *tUrns* table is a set of narnes of each [ibn's stars. The *radius* nlethod returns a circle centered at its first argulllent with radius equal to its second argurnent. The overlaps rnethod tests for spatial overlap. Nowshowing and Filrns are joined by the equijoin clause, \vhile Nowshowing and Countries are joined by the spatial overlap clause. The selections to 'Andorra' and fiInls containing 'Herbert the vVorrn' cornplete the query.

rrhcse two object-relational queries are sirnilar to SQL-92 queries but have SOIIIC unusual features:

- User-Defined Methods: User-defined abstract types are rnanipulated via their 1nethods, for exalnple, *is_herbert* (Section 23.2).
- Operators for Structured Types: A.long with the structured types available in the data rnodel, ()R,DBMSs provide the natural Inethods for those types. For exarnple, the ARRAY type supports the standard array

operation of accessing an array elenlent by specifying the index; *F.stars[l]* returns the first elernent of the array in the *staTs* cohllnn of film *F* (Section 23.3).

- Operators for Reference Types: Reference types are *dereferenced* via. an arrow (–> notation (Section 23.6.2).

To suuullarize the points highlighted by our 1110tivating exanlple, traditional relational systenls offer liInited flexibility in the data types available. Data is stored in tables and the type of each field value is lirnited to a siulple atornic type (e.g., integer or string), with a sl11all, fixed set of such types to choose frarn. This lirnited type systern can be extended in three Inain ways: user-defined abstract data types, structured types, and reference types. Collectively, we refer to these new types as complex types. In the rest of this chapter, we consider how a DBMS can be extended to provide support for defining new complex types and rnanipulating objects of these new types.

23.2 STRUCTURED DATA TYPES

SQL:1999 allows users to define new data types, in addition to the built-in types (e.g., integers). In Section 5.7.2, we discussed the definition of new *distinct* types. Distinct types stay within the standard relational model, since values of these types rnust be atornic.

SQL:1999 also introduced two type constructors that allow us to define new types with internaJ structure. Types defined using type constructors are called structured types. This ta.kes us beyond the relational model, since field values need no longer be atornic:

- RDW(n1 t_1, ..., *nn* t_n): A type representing a row, or tuple, of *n* fields \vith fields $n_1, \ldots, n_n$ of types $t_1, \ldots, t_n$ respectively.
- base ARRAY [iJ): A type representing an array of (up to) i base-type iterns.

The theater_t type in Figure 23.1 illustrates the new ROW data type. In SQL:1999, the ROW type has a special role because every table is a collection of ro\vs -.·every table is a set of 1'o\vs or a. rnultiset of rc)\vs. Values of other types can appear only as field values.

The *stars* field of table Filrns illustrates the new ARRAY type. It is an array of upto 10 elements, each of \vhich is of type VARCHAR(25). Note that 10 is the rnaxirnurn nurnber of elements in the array; at any tiTne, the array (unlike, say,

SQL:1999 Structured Data Types: Several conunercial systems, including IBM DB2, Infonnix UDS, and Oracle 9i support the ROWand ARRAY constructors. The `listof`, `bagof`, and `setof` type constructors are :not included in SQL:1999. Nonetheless, commercial systerIls support sorne of these constructors to varying degrees. Oracle supports nested relations and arrays, but does not support fully cornposing these constructors. InfOI'mix supports the setof, hagof, and Ustof constructors and allows thern to be cornposed. Support in this area varies \videly across vendors.

in C) can contain fewer elenlcnts. Since SQL:1999 does not support rnultidiInensional arrays, *vector* rnight ha,ve been a rnore accura,te narne for the array constructor.

The power of type constructors cornes froIn the fact that they can be cornposed. The following row type conta,ins a field that is an array of at Inost 10 strings:

ROW(filrnno: integer, *stars:* VARCHAR(25) ARRAY [10])

The row type in SQL:1999 is quite general; its fields can be of any SQL:1999 data type. Unfortunately, the arra.y type is restricted; elernents of an array cannot be arrays thcrnselves. Therefore, the following definition is illegal:

(integer ARRAY [5]) ARRAY [10]

23.2.1 Collection Types

SQL:1999 supports only the ROW a,nel ARRAY type constructors. Other COUUllon type constructors include

- listof(base): A. type representing a sequence of base-type itcrllS.
- setof(base): A type representing a *set* of base-type HeIns. Sets cannot contain duplicate elements.
- bagof(base): A type representin.g a, *bag* or *multiset* of base-type iterns.

Types llsing `listof`, ARRAY, bagof, or `setof` as the outennost type constructor are sometimes referred to as collection types or bulk data types.

The lack of support for these collection types is recognized as a weakness of SQL:1999's support for cornplex objects and it is quite possible that SODle of these collection types will be added in future revisions of the SQL standard.[1]

23.3 OPERATIONS ON STRUCTURED DATA

rIhe I)131V18 provides built-in Inethods for the types defined using type constructors. These lnethods are analogous to built-in operations such as addition and rIlultiplication for atcnnic types such as integers. In this section we present the Illethods for various type constructors and illustrate ho\v SQL queries can create and rnanipulate values with structured types.

23.3.1 Operations on Rows

Given an iteul i whose type is **ROW**$(n_1\ t_1, ..., n_n\ t_n)$, the field extraction rnethod allo\vs us to access an individuaJ field n_k llsing the traditional clot notation $i.n_k$. If ro\v constructors are nested in a type definition, dots rnay be nested to access the fields of the nested row; for exarnple $i.n_k.m_l$. If we have a collection of rows, the dot notation gives us a collection as a result. :For exarnple, **if** i is a list of rows, $i.n_k$ gives us a list of itcrns of type t_n; if i is a set of rows, $i.n_k$ gives us a set of iterns of type t_n.

[This nested-dot notation is often called a path expression, because it describes a path through the nested structure.

23.3.2 ()peratiolls on Arrays

Array types support an 'array index' rnethod to allow 11sers to access array iterns at a, particular offset. A. postfix 'square bracket' syntax is usually used. Since the nuruber of elernents can vary, there is an operator (CARDINALITY) that returns the nU1nbe1' of elerIlents ıı1 the array. The varia,hle nurnl)er of elernents also rn.otivates an operator to C:Ollcatenate two arrays. The following exanlple illustrates these operations on SQL:1999 arrays.

```
SELECT  F.fillnIlo, (F.staTs || ['Brando', 'Pacino'])
FROM    FilrnsF
WHERE   CARDINALITY(F.stars) < 3 AND F.stars[1]='Redford'
```

1According to Jinl Melton, the editor of the SQL:1999 standard, these collection types were considered for inclusion but omitted because some problems with their specifications were discovered too late for correction in the SQL:1999 time-frame.

For each fibn with Redford as the first star[2] and fewer than three stars, the result of the query contains the film's array of stars concatenated with the array containing the two elcrnents 'Brando' and 'Pacino'. Observe how a value of type array (containing Brando and Pacino) is constructed through the use of square brackets in the SELECT clause.

23.3.3 Operations on Other Collection Types

Although only arrays are supported in SQL:1999, future versions of SQL are expected to support other collection types, and we consider what operations are appropriate over these types of data. provide such operations. Our discussion is illustrative and not Ineant to be cOlnprehensive. For exarnple, one could additionally allow aggregate operators *count*, *sum*, *avg*, *max*, and *rnin* to be applied to any object of a collection type with an appropriate base type (e.g., INTEGER). ()ne could also support operators for type conversions. For exaInple, one could provide operators to convert a rnultiset object to a set object by elirninating duplicates.

Sets and Multisets

Set objects can be cornpared using the traditional set methods $\subset, \subseteq, =, \supseteq, \supset$. An iteln of type setof(faa) can be cornpared with an iteln of type faa using the $\in$ rnethod, as illustrated in Figure 23.3, which contains the cornparison *'Herbert the Worm'* $\in$ *F.stars*. T\vo set objects (having elernents of the saIne type) can be cornbined to forIlI a new object using the $\cup$, $\cap$, and -- operators.

Each of the Inethods for sets can be defined for Inultisets, taking the nUlnber of copies of elernents into account. The $\cup$ operation simply adds up the nurnber of copies of an elernent, the $\cap$ operation counts the lesser nU1nbel' of tirnes a given elernent appears in the two input rnultisets, and $-$ subtracts the nurnber of ti1nes a given e1ernent appears in the second lnultiset frorn the nUlnber of tinles it appears in the first Inultiset. For exarnplc, using rnultiset scrnantics $\cup$ ({1,2,2 2}, {2,2,3}) = {I,2,2,2,2,2,3}, $\cap$ ({"1.,2,2,2"}, {2,2,3}) = {2,2}; and I · ({1,2,2,2}, {2,2,3}) = {1,2}.

Lists

Traditional list operations include *head*, \vhich returns the first ele1nent; *tail*, which returns the list obtained by rCIIloving the first elernent; *prepend*, which

2Note that the first element in an SQL array has index value 1 (not 0, as in sorne languages).

takes an elcrnent and inserts it as the first elernent in a list; and *append*, which appends one list to another.

23.3.4 Queries Over Nested Collections

We no\v present SOHle exar11ples to illustrate ho\v relations that contain nested collections can be queried, using SQL syntax. In particular, extensions of the relational rnodel with nested sets and rnultisets have been \videly studied and \ve focus on these collection types.

We consider a variant of the FihIIS relation from Figure 23.1 in this section, with the *stars* field defined as a setof (VARCHAR[25]), rather than an array. Each tuple describes a filrn, uniquely identified by *filrnno,* and contains a set (of stars in the filrn) as a field value.

Our first exarnple illustrates how we can apply an aggregate operator to such a nested set. It identifies filrns with r11or8 than two stars by counting the nurnber of stars; the CARDINALITY operator is applied once per FilnIs tuple. [3]

```
SELECT  F.filmno
FROM    Filrns F
WHERE   CARDINALITY(F.stars) > 2
```

Our second query illustrates an operation called unnesting. Consider the instance of Filrns shown in Figure 23.4; we have olnitted the *director* and *budget* fields (included in the Filnls schema in Figure 23.1) for simplicity. A flat version of the saIne inforrna.tion is shown in Figure 23.5; for each filrn and star in the £ibn, we have a tuple in Filrns_flat.

filmno	*title*	*stars*
98	Casablanca	{Bogart, Bergluan}
54	Earth vVorrns Are Juicy	{Herbert, Wanda}

Figure **23.4** A Nested Relation, Films

1"11e follc)\ving query generates the instance of Films_flat from Fihns:

```
SELECT  FJilrnno, F.title, S AS star
FROM    FilrnsF, F.stars AS S
```

[3] SQL:1999 does not support set or rnultiset values, as we noted earlier. If it did, it would be natural to allow the CARDINALITY operator to be applied to a set-value to count the nUluber of elements; we have used the operator in this spirit.

filmno	*title*	*star*
98	Casablanca	Bogart
98	Casablanca	Bergman
54	Earth Worms Are Juicy	Herbert
54	Earth Worms Are Juicy	Wanda

"Figure 23.5 A Flat Version, Films_flat

The variable F is successively bound to tUI>les in Filrns, and for each value of F, the vaTiable S is successively bound to the set in the *stars* field of F. Conversely, we Inay want to generate the instance of F'ilrns frorn FilIns_fiat. We can generate the Filrns instance using a, generalized fonn of SQL's GROUP BY C011struct, as the following query illustrates:

```
SELECT    F.filmno, F.title, set_gen(F.star)
FROM      Fihns.Jlat F
GROUP BY  F.fihnno, F.title
```

This oxaluplc introduces a new operator *set_gen*, to be used with GROUP **BY**, that requires sorne explanation. The GROUP BY clause partitions the Films_flat table by sorting on the *filmno* attribute; all tuples in a given partition have the sanle *filrnno* (and therefore the sarne *title*). Consider the set of values in the *star* cohunn of a given partition. In an SQL-92 query, this set rnust be surnluarized by applying an aggregate operator such as COUNT. Now that we allow relations to contain sets as field values, however, we can return the set of *star* values as a field value in a single answer tuple; the answer tuple also contains the *filmno* of the corresponding partition. rrhe *set_gen* operator collects the set of *star* values in a paTtition and creates a set-valued object. This operation is called nesting. We can irnagine similar generator functions for creating Inuitise1's, lists, and so on. However, such generators are not included in SQL:1999.

23.4 ENCAPSULATION AND ADTS

Consicler the Frames table of Figure 23.1. It has a colun111 *image* of type `jpeg_image`, which stores a cOlnpressed image representing a single frarne of a film. The jpeg_image tYI)C is not one of the DBMS's l>uilt-in types and was defined 1)y a user for the I)inky application to store ilna,ge data cornpressed using the JPEG stanclard. As another exarllple, the Countries table defined in Line 7 of Figure 23.1 has a colurnn *boundary* of t,ype polygon, which contains representations of the shapes of countries' outlines on a world rnap.

.Allowing users to define arbitrary new data types is a key feature of ORDBMSs. The DBMS allows users to store and retrieve objects of type `jpeg_image`, just like an object of any other type, such as integer. New atomic data types usually need to have type-specific operations defined by the user who creates thern. For example, one rnight define operations on an irnage data type such as compress, rotate, shrink, and crop. The ccnnbination of an atolIlic data type and its associated rnethods is called an abstract data type, or ,A,DT. Traditional SQL COlnes with built-in ADTs, such as integers (with the associated arithnletic rnethods) or strings (with the equality, cornparison, and LIKE lllethods). Object-relational systerns include these ADT's and also allow users to define their o\vn ADTs.

The label *abstract* is applied to these data types because the database systerIl does not need to know how an ADT's data is stored nor ho\v the ADT's rnethods work. It rnerely needs to know what rnethods are availa,ble and the input and output types for the rnethods. I-Eding ADT internals is called encapsulation.[4] Note that even in a. relational systern, atolnic types such as integers have associated rnethods that encapsulate the1n. In the case of integers, the standard Inethods for the ADT are the usual aritlunetic operators and COIl1parators. To evaluate the addition operator on integers, the database systenl need not understand the laws of addition it l11erely needs to know how to invoke the addition operator's code and what type of data to expect in return.

In an object-relational systenl, the Silllplification due to encapsulation is critical because it hides any substantive distinctions between data types and allows an OR,DB1VIS to be iInplernented \vithout anticipating the types and rnethods that users Inight want to add. For exarnple, (l,dding integers and overlaying irnages can be treated unifonnly by the systern, with the only significant distinctions being that different code is invoked for the two operations and differently typed objects are expected to be returned froIll that code.

23.4.1 Defining Methods

To register a new rnethod for a user-defined data type, users rnust write the code for the nlcthod and then infor1n the database systcrI1 about the Inethod. The cod.e to be written depends on the languages supported by the DBlVIS and, possibly, the operating systerH in question. For example, the ORDBMS Inay handle Java code in the Linux operating systern. In this case, the lnet,hod code nlu,st be written iII Java and cOlnpiled into a Java bytecode file stored in. a Linux file system. 'Then an SQL-style luethod registration eOllunand is given to the ()I:ll)Bl\VlS so that it recognizes the new rnethod:

[4]Some ORDBMSs actually refer to ADTs as opaque types because they are encapsulated and hence one cannot see their details.

Packaged ORDBMS Extensions: Developing a set of user-defined types and rnethods for a particular application·····-sayimage management— can involve a significant aIIlount of work and dornain-speeific expertise. As a result, most ORDBMS vendors partner with third parties to sell prepackaged sets of ADrrs for particular domains. Inforn1ix calls these extensions *DataBlades,* Oracle calls theln *Data Cartridges,* IBM calls thern *DB2 Extenders,* and so on. These packages include the ADT 11lethod code, DDL scripts to automate loading the ADTs into the system, and in some cases specialized access methods for the data type. Packaged ADT extensions are analogous to the class libraries available for object-oriented programIning languages: They provide a set of objects that together address a COlnnlon task.

SQL:1999 has an extension called SQL/MIVI that consists of several independent parts, each of which specifies a type library for a particular kind of data. The SQL/MM parts for Full-Text, Spatial, Still Iillage, and Data Mining are available, or nearing publication.

```
CREATE FUNCTION is_sunrise(jpeg_image) RETURNS boolean
        AS EXTERNAL NAME '/a/b/c/dinky.class' LANGUAGE 'java';
```

This statenlent defines the salient aspects of the lllethod: the type of the associated ADT, the return type, and the location of the code. Once the method is registered, the DBMS uses a Java, virtual lnachine to execute the code[5]. Figure 23.6 presents a nUlnber of rnethod registration cOllllnands for our Dinky database.

```
1. CREATE FUNCTION thumbnail(jpeg_image) RETURNS jpeg_image
        AS EXTERNAL NAME '/a/b/c/dinky.class' LANGUAGE 'java';
2. CREATE FUNCTION is_sunrise(jpeg_image) RETURNS boolean
        AS EXTERNAL NAME '/a/b/e/dinky.class' LANGUAGE 'java';
3. CREATE FUNCTION isJnerbert(jpeg_image) RETURNS boolean
        AS EXTERNAL NAME '/a/b/c/dinky.class' LANGUAGE 'java';
4. CREATE FUNCTION radius(polygon, float) RETURNS polygon
        AS EXTERNAL NAME '/a/b/c/dinky.class' LANGUAGE 'java';
5. CREATE FUNCTION overlaps(polygon, polygon) RETURNS boolean
        AS EXTERNAL NAME '/a/b/c/dinky.class' LANGUAGE 'java';
```

Figure 23.6 Method Registration Conunands for the Dinky Database

51n the case of non-portable cOll1piled code written, for example, in a language like C++"—···the D131v18 uses the operating system's dynamic linking facility to link the method code into the database system so that it can be invoked.

rrype definition statelnents for the user-defined atornic data types in the Dinky scherna are given in Figure 23.7.

1. CREATE ABSTRACT DATA TYPE jpeg_image
   (*internallength* = VARIABLE, *input* = jpeg_in, *output* = jpeg_out);
2. CREATE ABSTRACT DATA TYPE polygon
   (internallength = VARIABLE, *input* = polyjn, *output* = poly_out);

Figure 23.7 Atomic Type Declaration Commands for Dinky Database

23.5 INHERITANCE

We considered the concept of inheritance in the context of the ER, model in Chapter 2 and discussed how ER diagrarns with inheritance 'were translated into tables. In object-database systems, unlike relational systerns, inheritance is supported directly and allows type definitions to be reused and refined very easily. It can be very helpful when modeling similar but slightly different classes of objects. In object-database systerns, inheritance can be used in two ways: for reusing and refining types and for creating hierarchies of collections of sirnilar but not identical objects.

23.5.1 Defining Types with Inheritance

In the Dinky database, we rnodel rnovie theaters with the type theater_t. Dinky also wants their database to represent a new rnarketing technique in the theater business: the *theater-cafe,* which serves pizza and other rneals while screening movies. rrheater-cafes require additional inforrnation to be represented in the database. In particular, a theater-cafe is just like a theater, but has an additional attribute representing the theater's IIlenu. Inheritance allows us to capture this 'specialization' explicitly in the database design with the followiIlg DDL staternent:

```
CREATE TYPE theatercafe_t UNDER theater_t (rn,enu text);
```

This staternent creates a new type, theatercafe_t, which has the sarne attributes and rnethods as theater_t, plus one additional attribute *menu* of type text. Methods defined on theater_t apply to objeets of type theatercafe_t, but not viee versa. We say that theatercafe_t inherits the attributes and rnethods of theater_t.

Note that the illherita,nce rnechanisrll is not rnerely a rnacro to shorten CREATE staternents. It creates an explicit relationship in the database between the subtype (theatercafe_t) and the supertype (theater_t*):An object of the*

subtype is also considered to be an object of the supertype. This treatment Ineans that an.y operations that apply to the supertype (nlcthods as well as query operators, such as projection or join) also apply to the subtype. This is generall.y· expressed in the follo\ving principle:

> The Substitution Principle: Given a supertype *A* and a subtype *B*, it is always possible to substitute an object of type *B* into a legal expression written for objects of type *A*, without producing type errors.

This principle enables easy code reuse because queries and Inethodswritten for the supertype can be applied to the subtype without I1lodification.

Note that inheritance can also be used for atomic types, in addition to ro\v types. Given a supertype image_t with rnethods *title()*, *number_of_colors()*, and *d'isplay()*, we can define a subtype thumbnail_image_t for slllall irnages that inherits the rnethods of image_to

23.5.2 Binding Methods

In defining a subtype, it is sornetiInes useful to replace a rnethod for the sII-pertype with a new version that operates differently on the subtype. Consider the image_t type and the subtype jpeg_image_t frorH the Dinky database. lJnfortunately, the *display()* rnethod for standard images does not work for JPEG irnages, which are specially cOlnpressed. Therefore, in creating type jpeg_image_t, we write a special *display()* rnethod for JPEG iruages and register it with the database systern using the CREATE FUNCTION cOIIuuand:

```
CREATE FUNCTION display(jpeg image) RETURNS jpeg_image
        AS EXTERNAL NAME '/a/b/c/jpeg.class' LANGUAGE 'java.';
```

Registering a new rnethod with the sarne name as an old rnethod is called overloading the luethod narne.

Because of overloading, the systern Inust understand 'which rnethod is intended in a particular expression. For example, when the systern needs to invoke the *display*() rnethod on an object of type jpeg_image_t, it uses the specialized *display* rnethocL When it needs to invoke *display* on an object of type image_t that is not otherwise subtyped, it invokes the standard *display* Inethod. The process of deciding which rnethod to invoke is called binding the rnethod to the object. In certain situations, this binding can be done when an expression is parsed (early binding), but in other cases the 1n08t specific type of an object cannot be known until rl.ln-tinle, so the rnethod cannot be l)ound until then (late binding). Late birlding fa,cilties acId flexibility but can rnake it harder

for the user to reason about the Inethods that get invoked for a given query expreSSIon.

23.5.3 Collection Hierarchies

Type inheritance was invented for object-oriented progranuning languages, and our discussion of inheritance up to this point differs little £roln the discussion one Inight find in a book on an object-oriented language such as C++ or Java.

However, because database systerns provide query languages over tabular data sets, the lnechanisnls fronl prograrnning languages are enhanced in object databases to deal with tables and queries as well. In particular, in objeet-relational systelIls, we can define a table containing objects of a particular type, such as the Theaters table in the Dinky sehenla. Given a new subtype, such as theatercafe_t, we would like to create another table Theater_cafes to store the inforrnation about theater cafes. But, when writing a query over the Theaters table, it is sornetirnes desirable to ask the saIne query over the rrheater_cafes table; after all, if we project out the additional C01UIIII1S, an instance of the Theater_cafes table can be regarded as an instance of the Theaters table.

R,ather than requiring the user to specify a separate query for each such table, we can infonn the systern that a new table of the subtype is to be treated as part of a table of the supertype, with respect to queries over the latter table. In our exalnple, we can say

```
CREATE TABLE Thea,ter_Cafes OF TYPE theatercafe_t UNDER Theaters;
```

This staternent tells the systern that queries over the Theaters table should actually be run over all tuples in both the rrheaters and rrheater_Cafes tables. In such cases, if the subtype definition involves rnethod overloading, late-binding is used to ensure that the appropriate rnethods are called for each tuple.

In general, t11.e UNDER clause can be used to genera,te an arbitrary tree of tables, called (a collection hierarchy. Queries over a particular tal)le T in the hierarchy are run over all tuples in rr and its descendants. Sornetirnes, a user rnaywant the query to nUI only on rr and not on the descendants; additiona1 syntax, for exaInple, the key\vord ONLY, can be used in the query's FROM clause to achieve this effect.

23.6 OBJECTS, OIDS, AND REFERENCE TYPES

In object-elatabase systerns, data objects can be given an object identifier (oid), \vhich is sotne value that is unique in the database across tirne. The

OIDs: IBM DB2, Informix lJDS, and Oracle 9i support REF types.

DBMS is responsible for generating aids and ensuring that an oid identifies an object uniquely over its entire lifetime. In SOHle systenls, all tuples stored in any table are objects and autornatically assigned unique oids; in other systenls, a user can specify the tables for 'which the tuples are to be assigned aids. Often, there are also facilities for generating oids for larger structures (e.g., tables) as well as slnaller structures (e.g., instances of data values such as a copy of the integer 5 or a .TPEG image).

An object's aid can be used to refer to it from elsewhere in the data. An oid has a type similar to the type of a pointer in a progralnll1ing language.

In SQL:1999 every tuple in a table can be given an aid by defining the table in ternlS of a structured type and declaring that a REF type is associated with it, as in the definition of the Theaters table in Line 4 of Figure 23.1. Contrast this with the definition of the Countries table in Line 7; Countries tuples do not have associated aids. (SQL:1999 also assigns 'oids' to large objects: This is the locator for the object.)

REF types have values that are unique identifiers or aids. SQL:1999 requires that a given REF type must be associated with a specific table. For exalnple, Line 5 of Figure 23.1 defines a cohllnn *theater* of type REF(theater_t). The SCOPE clause specifies that iterns in this colurnn are references to rows in the rrheaters table, which is defined in Line 4.

23.6.1 Notions of Equality

The distinction between reference types and reference-free structured types raises another issue: the definition of equality. Two objects having the saIne type are defined to be deep equal if and only if

1. The objects <1,1'e of atolnic type and have the same value.
2. The objects are of reference type and the *deep equals* operator is true for the two referenced objects.
3. The objects are of structured type and the *deep equals* operator is true for all the corresponding subparts of the two objects.

Two objects that have the same reference type are defined to be shallow equal if both refer to the saIne object (i.e., both references use the saUle aid). The

definition of shallow equality can be extended to objects of arbitrary type by taking the definition of deep equality and replacing *deep equals* by *shallow equals* in parts (2) and (3).

As an example, consider the cornplex objects ROW(538, *tS9,* 6-3-97, 8-7-97) and ROW(538, *i33,* 6-3-97, 8-7-97), whose type is the type of rows in the table Nowshowing (Line 5 of Figure 23.1). 1'hese two objects are not shallow equal because they differ in the second attribute value. Nonetheless, they rnight be deep equal, if, for instance, the oids *t89* and *t33* refer to objects of type theater_t that have the saIne value; for exarnple, tuple(54, 'Majestic', '115 King', '2556698').

While two deep equal objects Inay not be shallow equal, as the exarnple illustrates, two shallow equal objects are always deep equal, of course. 'The default choice of deep versus shallow equality for reference types is different across systenls, although typically we are given syntax to specify either semantics.

23.6.2 Dereferencing Reference Types

An item of reference type REF(basetype) is not the sarne as the basetype itenl to which it points. To access the referenced basetype itenl, a built-in deref() rnethod is provided along with the REF type constructor. For example, given a tuple from the Nowshowing table, one can access the *name* field of the referenced theater_t object with the syntax Nowshowing.deref*(theater).narne.* Since references to tuple types are comInon, SQL:1999 uses a Java-style arrow operator, which cOD.lbines a postfix version of the dereference operator with a tuple-type dot operator. The narne of the referenced theater can be accessed with the equivalent syntax *Nowshowing.theater-> narne,* as in Figure 23.3.

At this point we have covered all the basic type extensions used in the Dinky scherna in Figure 23.1. The reader is invited to revisit the scherna and exarnine the structure and content of each table and how the new features are used in the various sarnple queries.

23.6.3 URLs and DIDs in SQL:1999

It is instructive to note the differences between Internet lJRIJs and the oids in object systerns. First, oids uniquely identify a single object over all tirne (at least, until the object is deleted, when the oid is undefined), whereas the Web resource pointed at by an URL can change over tirue. Second, oids are sirnply identifiers and carry no physical infonnation about the objects they identify this rnakes it possible to change the storage location of an object without rnodifying pointers to the object. In contrast, URLs include network

addresses and often file-syst;enl names as well, meaning that if the resource identified by the URL has to move to another file or network address, then all links to that resource are either incorrect or require a 'forwarding' rnechanisH1. Third, oids are automatically generated by the DBMS for each object, whereas URLs are user-generated. Since users generate URLs, they often ernbed scrnantic inforlllation into the URL via rnachine, directory, or file names; this can becoine confusing if the object's properties change over tilne.

For URLs, deletions can be troublesorne: This leads to the notorious '404 Page Not Found' error. For oids, SQL:1999 allows us to say REFERENCES ARE CHECKED as part of the SCOPE clause and choose one of several actiol1swhen a referenced object is deleted. This is a direct extension of referential integrity that covers oids.

23.7 DATABASE DESIGN FOR AN ORDBMS

The rich variety of data types in an ORDBMS offers a database designer Inany opportunities for a rnore natural or lllore efficient design. In this section we illustrate the differences between RDBMS and ()RI)BMS database design through several exarnples.

23.7.1 Collection Types and ADTs

()ur first exarnple involves several space probes, each of which continuously records a video. A single video strearll is associated with each probe, and while this strearn was conected over a certail1 tiule period, we assurne that it is now a cOlllplete object associated with the probe. During the tirne period over which the video was collected, the probe's locatiol1\vas periodieaJly recorded (such infonnation can easily be piggy-backed onto the header portion of a video streanl conforrning to the MPEG sta,ndard). The inforrnation associated with a probe has three parts: (1) a *probe ID* that identifies a probe uniquely, (2) a *video stream*, and (3) a *location sequence* of $\langle time, location \rangle$ pairs. What kind of a database scherna should we use to store this infonnation?

An RDBMS Database Design

In an RDBMS, we rnust store each video strcanl as a BLO13 and each location sequellce as tuples in a table. A possible RDBMS database design follo\vs:

Probes(*pid:* integer, *time:* timestamp, *lat:* real, *long:* real,
camera: string, *video:* BLOB)

There is a single table caned Probes and it has several rows for each probe. Each of these rows has the same *pid*, *camera* and *video* values, but different *time*, *tat*, and *long* values. (We have used latitude and longitude to denote location.) The key for this table can be represented as a functional dependency: *IJTLN* $\rightarrow$ *CV*, where *N* stands for longitude. There is another dependency: $P \rightarrow CV$. This relation is therefore not in BCNF; indeed, it is not even in 3NF. We ca:n decolupose Probes to obtain a BCNF scherna:

Probes_Loc(*pid:* integer, *time:* timestamp, *tat:* real, *long:* real)
Probes_Video(*pid:* integer, *camera:* `string`, *video:* BLOB)

This design is about the best we can achieve in an RDBMS. However, it suffers frorn several drawbacks.

First, representing videos as BLOBs lI1eanS that we have to write application code in an external language to lnanipulate a video object in the database. Consider this query: "For probe 10, display the video recorded between 1:10 P.M. and 1:15 P.M. on May 10 1996." We lnust retrieve the entire video object associated "lith probe 10, recorded over several hours, to display a segrnent recorded over five rninutes.

Next, the fact that each probe has an associated sequence of location readings is obscured, and the sequence inforrnatiol1 associated with a probe is dispersed across several tuples. A third drawback is that we are forced to separate the video infonnation froTn the sequence inforrnation for a probe. These lirnitations are exposed by queries that require us to consider all the infonnation associated with each probe; for example, "For each probe, print the earliest tirne at which it recorded, and the camera type." T'his query now involves a join of Probes_Loc and Probes_Video on the *pid* field.

An ORDBMS Database Design

;\n ORDBMS supports a lnuch better solution. First. we can store the video as an A.DT object and write rnethods that capture a.ny special rna,nipulation we wish to perforrr1. Second. because we are allowed to store structured types such as lists, we (:a,n stc)re the location sequence for a probe in a single tuple, along\vith thę video infonnation. This layout eliminates the need for joins in queries that involve both the sequence and video inforrnation. An ORDBMS design for our example consists of a single relation called Probes_AllInfo:

Probes_AllInfo(*pid:* integer, *locseq:* `location_seq`, *camera:* `string`; *video:* `mpeg_stream`)

This definition involves two new types, location_seq and mpeg_stream. The mpeg_stream type is defined as an ADT, with a Inethod *display()* that takes a start time and an end tirne and displays the portion of the video recorded during that interval. This rnethod can be irnplernented efficiently by looking at the total recording duration and the total length of the video and interpolating to extract the segnlcnt recorded during the interval specified in the query.

Our first query in extended SQL using this *display* lnethod follows. We now retrieve only the required segment of the video rather than the entire video.

```
SELECT display(P.video, 1:10 Po M May 10 1996, 1:15 Po M May 10 1996)
FROM   Probes_AllInfo P
WHERE  Popid = 10
```

Now consider the location_seq type. We could define it as a list type, containing a list of ROW type objects:

```
CREATE TYPE location_seq listof
        (row (time: timestamp, lat: real, long: real))
```

Consider the *locseq* field in a row for a given probe. This field contains a list of rows, each of which has three fields. If the ORDBMS implements collection types in their full generality, we should be able to extract the *time* colurnn from this list to obtain a list of timestamp values and apply the MIN aggregate operator to this list to find the earliest time at which the given probe recorded. Such support for collection types would enable us to express our second query thus:

```
SELECT   P.piel, MIN(P.locseq.tirne)
FROM     Probes._AllInfo P
```

Current ORDBMSs are not as general and clean as this exalnple query suggests. For instance, the systern rnay not recognize that projecting the *time* colurnn frorn a list of rows gives us a list of tirnestarnp values; or the systcru rnay allow us to apply an aggregate operator only to a table and not to a nested list value.

Continuing with our example, we lnay want to do specialized operations on our location sequences that go beyond the standard aggregate operators. For instance, we rnay want to define a lnethod that takes a tirne interval and COIIIputes the distance traveled by the probe during this interval. The code for this rnethod rnust understand details of a probe's trajectory and geospatial coordinate systenls. Fbl' these reasons, we rnight choose to define location_seq as an ADT\.

Clearly, an (ideal) ORDBMS gives us IllalIY useful design options that are not available in an RDBMS.

23.7.2 Object Identity

We now discuss S0l11e of the consequences of using reference types or aids. The use of aids is especially significant when the size of the object is large, either because it is a structured data type or because it is a big object such as an image.

Although reference types and structured types seem sirnilar, they are actually quite different. For example, consider a structured type my_theater tuple *(ina* integer, *name* text, *address* text, *phone* text) and the reference type theater ref (theater_t) of Figure 23.1. rrhere are irnportant differences in the way that database updates affect these two types:

- Deletion: Objects with references can be affected by the deletion of objects that they reference, while reference-free structured objects are not affected by deletion of other objects. For exaluple, if the Theaters table were dropped from the database, an object of type theater might change value to *null*, because the theater_t object it refers to has been deleted, while a similar object of type my_theater would not change value.
- Update: Objects of reference types change value if the referenced object is updated. Objects of reference-free structured types change value only if updated directly.
- Sharing versus Copying: An identified object can be referenced by llluitiple reference-type iterIls, so that each update to the object is reflected in IYlany places. To get a sirnilar effect in reference-free types requires updating all 'copies' of an object.

There are also irnportant storage distinctions between reference types and non-reference types, which rnight affect perfoI'rnance:

- Storage Overhead: Storing copies of a large value in rnultiple structured type objects IYlay use lnnch mo1'e space than stori.ng the value once and referring to it elsewhere through reference type objects. This additional storage requirelnent can affect both disk usage and buffer lnanagcrncnt (if IIlal1Y copies are accessed at once).
- Clustering: The subparts of a structured object are typically stored together on disk. Objects with references rna,Y point to other objects that are far a:way on the disk, and the disk arm Inay require significant rnOVClnent

OIDs and Referential Integrity: In SQL:1999, all the oids that appear in a cohunn of a relation are required to reference the same target relation. This 'scoping' makes it possil)]e to check oid refere:nces for 'referential integrity' just like foreign key references are checked. While current ORDBMS products supporting oids do not support such checks, it is likely that they will in future releases. This will nlake it rnnch safer to use aids.

to asserrlble the object and its references together. Structured objects can thus be l'nor8 efficient than reference types if they are typically accessed in their entirety.

Many of these issues also arise in traditional prograunuing languages such as C or Pascal, which distinguish between the notions of referring to objects *by value* and *by reference*. In database design, the choice between using a structured type or a reference type typically includes consideration of the storage costs, clustering issues, and the effect of updates.

Object Identity versus Foreign Keys

lJsing an oid to refer to an object is silnila,r to using a foreign key to refer to a tuple in another relation but not quite the same: An oid can point to an object of theater_t that is stored *anywhere* in the database, even in a field, whereas a foreign key reference is constrained to point to an object in a particular referenced relation. This restriction rnakes it possible for the DBMS to provide lnuch greater support for referential integrity than for arbitra,ry aid pointers. In general, if an object is deleted while there are still oid-pointers to it, the best the DBl\IIS can do is to recognize the situation by rnaintajning a reference count. (Even this lirnited support becornes irnpossible if oids can be copied freely.) Therefore, the responsibility for avoiding dangling references rests largely with the user if oids are llsed to refer to objects. This burdensolllc responsibility suggests that we should use oids with great ca.ution and use foreign keys instead \vhenever possible.

23.7.3 Extending the ER Model

The ER rnodel, as described in Chapter 2, is not adequate for ORDBMS design. We have to use an extended ER rnodel that supports structured attributes (i.e., sets, lists, arra,Ys as attribute values), distinguishes whether entities have ol)ject ids, and allows us to Inodel entities whose attributes include rnethods. We illustrate these connnents using an extended ER diagrarn to describe the

space probe data in Figure 23.8; our notational conventions are ad hoc and only for illustrative purposes.

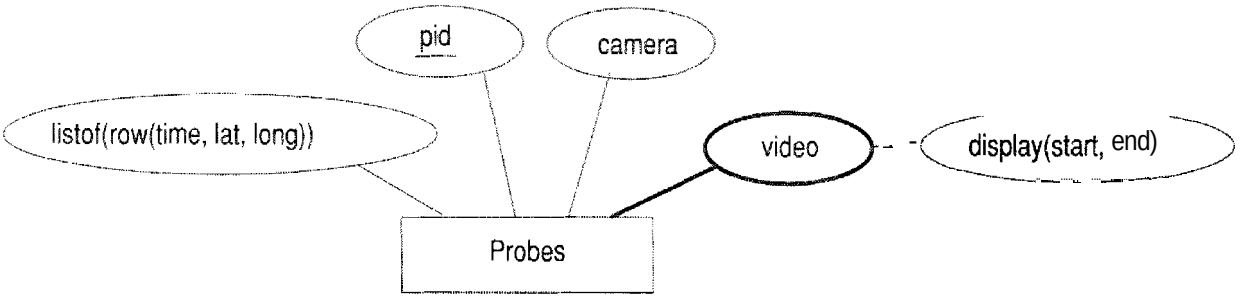


Figure 23.8 The Space Probe Entity Set

The definition of Probes in Figure 23.8 has two new aspects. First, it has a structured-type attrilnlte **listof** (row(*time, lat, lo'ng)); each* value assigned to this attribute in a Probes entity is a list of tuples with three fields. Second, Probes has an attribute called *video* that is an abstract data type object, \vhich is iIldicatecl by a dark oval for this attribute with a dark line connecting it to Probes. Further, this attribute has an 'attribute' of its own, which is a rnethod of the *J\DT.*

Alternatively, we could rnodel each video as an entity by using an entity set called Videos. The association between Probes entities and Videos entities could then be captured by defining a relationship set that links thenl. Since each video is collected by precisely one probe and every video is collected by se)lne probe, this relationship can be rna,intained by simply storing a reference to a probe object with each Videos entity; this technique is essentially the second translation approach frorn ER, diagrams to tables discussed in Section 2.4.1.

If we also rnake Videos a weak entity set in this alternative design, we can add a referential integrity constraint that causes a Videos entity to be deleted \vhen the corresponding Probes entity is deleted. More generally, this alternative design illustrates a strong sirnila1'ity between storing references to objects and foreign keys; the foreign key rnechanisT11 achieves the saIne effect as storing oids, but in a controlled lUanneI'. If oids are used. the user rnusi ensure that there are no dangling references when an object is deleted, with very little support froIll the DBMS.

Finally, we note tllat a significant extension to the ER rHodel is required to support the design of nested collections. For example, if a location sequence is rnodeled as an entity, and we want to define an attribute of Probes that contains a set of such entities, there is no way to do this\vithont extending the ER model. We do not disC1ISS this I)oint furtller at the level of ER diagrams, but consider an exaJnple next that illustrates when to use a nest,ed collection.

23.7.4 Using Nested Collections

Nested collections offer great rnodeling power but also raise difficult design decisions. Consider the following way to rnodel location sequences (other infonnation about probes is ornitted here to sirnplify the discussion):

Probesl(pid: integer, *locseq:* location_seq)

rrhis is a good choice if the irnportant queries in the workload require us to look at the location sequence for a particular probe, as in the query "For each probe, print the earliest tirne at which it recorded and the caluera type." On the other hand, consider a query that requires us to look at all location sequences: "Find the earliest time at which a recording exists for *laf;=5, long=90."* This query can be answered 1110re efficiently if the following scherna is used:

Probes2(*pid:* integer, *time:* timestamp, *tat:* real, *long:* real)

The choice of scherna Blust therefore be guided by the expected workload (as always). As another example, consider the following scherna:

Can_Teachl *(cid:* integer, *teacheTs:* setof*(ssn:* string), *sal:* integer)

If tuples in this table are to be interpreted as "Course *cid* can be taught by any of the teachers in the *teacheTs* field, at a cost *sal."* then we have the option of using the following schenla. instead:

CarLTeach2(*cid:* integer, *teachCT_8sn:* string, *sal:* integer)

A choice between these two alternatives can be Inade based on how we expect to query this table. On the other hand, suppose that tuples in CalL.Teachl are to be interpreted as "Course *cid* can be taught by the *tearnteacheT8,* at a cornbined cost of *sal."* CarLTeach2 is no longer a viable alternative. If we wanted to flatten Can_Teach1, we would have to use (1, separate table to encode tearns:

Can_Teach2(*cid:* integer, *team_id:* oid, *sal:* integer)
Teams(*tid:* oid, *ssn:* string)

As these exarnples illustrate, nested collections are appropriate in certain situations, but this fea,ture can easily be rnisused; nested collections should therefore be used ·with care.

23.8 ORDBMS IMPLEMENTATION CHALLENGES

The enhanced functionality of ORDBMSs raises several irnp1ernentation challenges. SOlne of these are 'well understood and solutions have been irnp1enlented in products; others are subjects of current research. In this section \ve exarnine a few of the key challenges that arise in irnplernenting an efficient, fully functional ORDBMS. Many rnore issues are involved than those discussed here; the interested reader is encouraged to revisit the previous chapters in this book and consider whether the irnplernentation techniques described there apply naturally to ORDBMSs or not.

23.8.1 Storage and Access Methods

Since object-relational databases store new types of data, ORDBMS implernentors need to revisit some of the storage and indexing issues discussed in earlier chapters. In particular, the system lllust efficiently store ADT objects and structured objects and provide efficient indexed access to both.

Storing Large ADT **and** Structured Type **Objects**

Large ADT objects and structured objects cornplicate the layout of data on disk. This problern is well understood and has been solved in essentially all ORDBMSs and OODBMSs. We present Sallie of the main issues here.

User-defined ADTs can be quite la,rge. In particular, they can be bigger than a single disk page. Large ADTs, like BLOBs, require special storage, typically in a different location on disk frorn the tuples that contain them. Disk-based pointers are rnaintained frorn the tuples to the objects they contain.

Structured objects can also be large, but unlike ADrr objects, they often vary in size during the lifetirne of a database. For exarnple, consider the *stars* attribute of the *film,s* table in Figure 23.1. As the years pass, SOlne of the 'bit actors' in an old rnovie may becorne famous.[6] When a bit actor hecornes farnous, Dinky rnight want to advertise his or her presence in the earlier filrns. This involves an insertion into the *stars* attribute of an individual tuple in *filrns*. Because these bulk attributes can grow arbitrarily, flexible disk layout rnechanisrns are required.

[6] A famous example is Marilyn Monroe, who had a bit part in the Bette Davis classic *All About Eve*.

An additional COlllplication arises \'lith array types. Traditionally, array elelnents are stored sequentially on disk in a row-by-row fashion; for example

$$A_{11}, \ldots, A_{1n}, A_{21}, \ldots, A_{2n}, \ldots, A_{m1}, \ldots, A_{mn}$$

However, queries rnay often request suba.rrays that are not stored contiguously on disk (e.g., $A_{11}, A_{21}, \ldots, A_{rn1}$). Such requests can result in a very high I/O cost for retrieving the subarray. To reduce the nurnber of l/Os required, arrays are often broken into contiguous *chunks*, which are then stored in some order on disk. Although each chunk is sorne contiguous region of the array, chunks need not be row-by-row or colurnn-by-colurllll. For exalnple, a chunk of size 4 111ight be $All, A_{12}, A_{21}, A_{22}$, 'which is a square region if we think of the array as being arranged row-by-row in two dimensions.

Indexing New Types

One ilnportant reason for users to place their data in a database is to allow for efficient access via indexes. Unfortunately, the standard RDBMS index structures support only equality conditions (B+ trees and hash indexes) and range conditions (B+ trees). An irnportant issue for OR,DB1\IISs is to provide efficient indexes for AD'I' rnethods and operators on structured objects.

Many specialized index structures have been proposed by researchers for particular applications such as cartography, genorne research, 11lultirnedia repositories, Web search, and so on. An ORDBMS cornpany cannot possibly inlplernent every index that has been invented. Instead, the set of index structures in an ORDBMS should be user-extensible. Extensibility would allow an expert in cartography, for exanlple, to not only register an ADT for points on a rnap (i.e., latitude··longitude pairs) but also irnplernent an index structure that supports natural rnap queries (e.g., the R-tree, \vhich lnatches cOllclitions such as "Find rne all theaters within 100 Iniles of Andorra"). (See Chapter 28 for 1110re on R-trees and other spatial indexes.)

One way to rnake the set. of index structures extensible is to publish an *access method interface* that lets users irnplcrnent an index structure *outside* the DBMS. The index and data can be stored in a file systeIll and the DBMS sirnply issues the *open*, *next*, and *close* iterator requests to the user's external index code. Such functionality rnakes it possible for a user to connect a I)Bl\1S to a Web search engine, for example. A rnain drawback of this approach is that data in an external index is 11ot protected by the])B1V1S'8 support for concurrency and recovery. 1\n alternative is for the ORDBMS to provide a generic 'template' index structure that is sufficiently general to encornpass rnost index structures that usersrn.ight invent. Because snell a structure is implemented within the DBMS, it can support high concurrency and recovery. The *Gener-*

alized Search Tree (GiS'r) is suell a structure. It is a ternplate index structure based on B+ trees, \vhich aJlo\vs IIIOSt of the tree index structures invented so far to be irnplernented'with only a few lines of user-definedAD'T' code.

23.8.2 Query Processing

ADTs and structured types call for l1e\v functionality in processing queries in ORDBMSs. They also change a nurnber of assumptions that affect the efficiency of queries. In this section we look at two functionality issues (u8er-defined aggregates and security) and two efficiency issues (rnethod caching and pointer swizzling).

User-Defined Aggregation Functions

Since users are allowed to define new rnethods for their ADTs, it is not unreasonable to expect thern to want to define new aggregation fUllctions for their ADTs as well. For example, the usual SQL aggregates·····---CDUNT, SUM, MIN, MAX, AVG--are not particularly appropriate for the image type in the Dinky schema.

Most ORDBMSs allow users to register new aggregation functions \vith the systern. To register an aggregation function, a user lnust iruplenlent three methods, which we call *initialize*, *iterate*, and *terrninate*. The *initialize* rnethod initializes the internal state for the aggregation. The *iterate* rnethod updates that state for every tuple seen, while the *terrninate* rnethod C0111putes the aggregation result based on the final state and then cleans up. As an exarnple, consider an aggregation function to cornpute the second-highest value in a field. The *initialize* call would allocate storage for the top two values, the *iterate* call would corupare the current tuple's value with the top two and update the top two as necessary, and the *terminate* call \vould delete the storage for the top two values, returning a copy of the second-highest value.

Method Security

ADTs give users the power to add code to the DBMS; this power can be abused. A buggy or rnalicious ADT rnethod can bring do\vn the database server or even corrupt the database. The DBMS lnust have rnechanisrns to prevent buggy or rnalicious user code from causing probleIlls. It 1nay rnake sense to overricle these rnechanislIls for efficiency in production environrnents with vendor-supplied rnethods. I-Io\vever, it is irnportant for the rnechanisrns to exist, if only to support delJugging of *J\DT* rnethocls; otherwise rnethod writers

\vould have to \vrite bug-free code before registering their rnethods with the DBMS--not a very forgiving prograllIuning environlnent.

()ne rnechanisrn to prevent problerns is to have the user rnethods be *interpreted* rather than *compiled.* The DBMS can check that the rnethod is well behaved either by restricting the power of the interpreted language or by ensuring that each step taken by a rnethod is safe before executing it. Typical interpreted la.nguages for this purpose include Java and the procedural portions of SQL:1999.

An alternative rnechanislll is to allow user methods to be cOlnpiled frorn a general-purpose progranuning language, such as C++, but to run those rneth-ods in a different address space than the DBMS. In this case, the DBMS sends explicit interprocess cOl1uIlunications (IPCs) to the user method, which sends IPCs back in return. This approach prevents bugs in the user methods (e.g., stray pointers) frorn corrupting the state of the DBNIS or database and prevents rnalicious methods frorn reading or Inodifying the DBMS state or database as well. Note that the user writing the method need not know that the DBMS is running the method in a separate process: The user code can be linked with a 'wrapper' that turns method invocations and return values into IPCs.

Method Caching

User-defined ADT methods can be very expensive to execute and can account for the bulk of the time spent in processing a query. During query processing, it may 11lake sense to cache the results of methods, in case they are invoked lllultiple times with the same arglunent. Within the scope of a single query, one can avoid calling a Inethod twice on duplicate values in a colurnn by either sorting the table on that colullln or using a hash-based scherne ruuch like that used for aggregation (see Section 14.6). An alternative is to rnaintain a *cache* of rnethod inputs and rnatching outputs as a table in the database. Then, to find the value of a rnethod on particular inputs, we essentially join the input tuples with the cache table. rrhese two approaches can also be cornbined.

Pointer Swizzling

In sorne applications, objects are retrieved into rnernory and accessed frequently through their oids; dereferencin.g rnust be irnplcrnented very efficiently. 801ne systerns rnaintain a ta,hle of oids of objects that are (currently) in InenlOI'.Y. \Vhen an object () is brought into memory, they check each oid contained in 0 and replace oids of in-memory objects by in-rncrl10ry pointers to those objects. This techrlique, caJled pointer swizzling, makes references to in-rnernory objects very .fast. rfhe downside is that when an object is paged out,

23.9 OODBMS

In the introduction of this chapter, we defined an OODBMS as a prograrnlIning language with support for persistent objects.vVhile this definition reflects the origins of OODBMSs accurately, and to a certain extent the irnplernentation focus of OODBMSs, the fact that OODBl\!ISs support *collection types* (see Section 23.2.1) rnakes it possible to provide a query language over collections. Indeed, a standard has been developed by the Object Database Management Group and is called Object Query Language.

OQL is sirnilar to SQL, with a SELECT----FROM--HWHERE---style syntax (even GROUP BY, HAVING, and ORDER BY are supported) and many of the proposed SQL:1999 extensions. Notably, OQL supports structured types, including sets, bags, arrays, and lists. The OQL treatrnent of collections is rnore uniforlll than SQL:1999 in that it does not give special treatrnent to collections of rows; for exalnple, ()QL allows the aggregate operation COUNT to be applied to a list to C0111pute the length of the list. OQL also supports reference types, path expressions,ADrrs and inheritance, type extents, and SQL-style nested queries. 1'here is also a standard Data Definition Language for OODB1\1Ss (Object Data Language, or ODL) that is sirnilar to the DDL subset of SQL but supports the additional features found in OODBMSs, such as ADT definitions.

23.9.1 The ODMG· Data Model and ODL

'The ODl\iIG data rnodel is the basis for an OODBl\iIS, just like the relational data 1nodel is the basis for an RDBMS. A database contains a collection of objects, which are sirnilar to entities in the ER rnode!. Every object has a unique aid, and a database contains collections of objects with Silllilar properties; such a collection is called a class.

The properties of a class are specified using ()l)L and are of three kinds: attributes, relationships, and rnethod8. _Attributes have an atolnic type or a structured type. ODl.l supports the **set**, bag, **list**, array, and **struct** t,ype constructors; these are just **setof**, bagof, **listof**, ARRAY, and ROW in the terrninology of Section 2:3.2.1.

R,elationships have a type that is either a reference to an object or a collection of such references. A relationship captures how an object is related to one or r1101'e objects of the same class or of a different class. A relationship in the ()IJIvIG- rnodel is really just a binary relationship in the sense of the ER Inodel. A rela.tionship has a corresponding inverse relationship; intuitively, it is the relationship 'in the other clirection.' For exarnple, if a Inovie is being

Class == Interface + Implenlentation: Properly speaking, a class consists of an interface together\vith an irnpleluentation of the interface. An ODL interface definition is implemented in an OODBMS by translating it into declarations of the object-oriented language (e.g., C++, Snlalltalk or Java) supported by the OODBMS. If we consider C++, for instance, there is a library of classes that irnplcrnent the ODL constructs. There is also an Object Manipulation Language (OML) specific to the programlning language (in our exanlple, C++), which specifies how database objects are manipulated in the progralnnl.ing .language. rr.he .goal is to seamlessly integrate the prograrnrning language and the database features.

shown at several theaters and each theater shows several rnovies, we have two relationships that are inverses of each other: *shownAt* is associated with the class of movies and is the set of theaters at which the given movie is being shown, and *nowShowing* is associated with the class of theaters and is the set of rnovies being shown at that theater.

Methods are functions that can be applied to objects of the class. There is no analog to methods in the ER or relational models.

The keyword interface is used to define a class. For each interface, we can declare an extent, which is the narne for the current set of objects of that class. The extent is analogous to the instance of a relation and the interface is analogous to the scherna. If the user does not anticipate the need to work with the set of objects of a given class-it is sufficient to manipulate individual objects—the extent declaration can be ornitted.

The following ()DL definitions of the Movie and Theater classes illustrate these concepts. (While these classes bear S(Hne resernblance to the Dinky database scherna, the reader should not look for an exact parallel, since we have rnodified the exarnple to highlight ()DL features.)

```
interface Movie
    (extent Movies key rnovieNarne)
    { attribute date start;
    attribute date end;
    attribute string rnovienarne;
    relationship Set⟨Theater⟩ shownAt inverse Theater::nowSho\ving;
    }
```

The collection of database objects whose class is Movie is called Movies. No two objects in lVIovies have the sarne *rnovieNarne* value, as the key declaration

indicates. Each lnovie is shown at a set of theaters and is shown during the specified period. (It \vould be rnore realistic to associate a different period with each theater, since a 1110vie is typically played at different theaters over different periods. While we can define a class that captures this detail, we have chosen a sirnpler definition for our discussion.) A theater is an object of class Theater, defined as:

```
interface Theater
    (extent Theaters key theaterNarne)
    { attribute string theaterName;
    attribute string address;
    attribute integer ticketPrice;
    relationship Set(Movie) nowShowing inverse .lVlovie::shownAt;
    float numshowingO raises(errorConntingMovies);
    }
```

Each theater shows several movies and charges the same ticket price for every movie. Observe that the *shownAt* relationship of Movie and the *nowShowing* relationship of Theater are declared to be inverses of each other. Theater also has a lllethod *numshowing()* that can be applied to a theater object to find the number of movies being shown at that theater.

ODL also allows us to specify inheritance hierarchies, as the following class definition illustrates:

```
interface SpecialShow extends lVlovie
    (extent SpecialShows)
    { attribute integer I11axinnunAttendees;
    attribute string benefitCharity;
    }
```

An object of class SpecialShow is an object of class Movie, with SOI11e additional properties, as discussed in Section 23.5.

23.9.2 OQL

rIhe ODMG query language OQL was deliberately designed to have syntax sirnilar to SQL to rnake it easy for users falniliar with SQL to learn ()QL. Let us begin with a query that finds pairs of Inovies and theaters such that the rnovie is shown at the theater and the theater is showing lHore than one rnovie:

```
SELECT Innarne: M.movieName, tnaIne: T.theaterName
```

```
FROM    lViovies M, lVLsho\vnAt T
WHERE   T.nurllshowing() > 1
```

The SELECT clause indicates how we can give Ilallles to fields in the result: The two result fields are called *rnnarne* and *tname*. The part of this query that differs frorn SQL is the FROM clause. The variable 1V[is bound in turn to each rnovie in the extent Movies. For a given rnovie M, we bind the variable T in turn to each theater in the collection $M.shownAt$. Thus, the use of the path expression $M.shownAt$ allows us to easily express a nested query. The following query illustrates the grouping construct in OQL:

```
SELECT    T.ticketPrice,
          avgNurn: AVG(SELECT P.T.nurnshowingO FROM partition P)
FROM      l'heaters T
GROUP BY  T.ticketPrice
```

For each ticket price, we create a group of theaters with that ticket price. This group of theaters is the partition for that ticket price, referred to using the OQL keyword partition. In the SELECT clause, for each ticket price, we cornpute the average nunlber of rnovies shown at theaters in the partition for that ticketPrice. OQL supports an interesting variation of the grouping operation that is missing in SQL:

```
SELECT    low, high,
          avgNllln: AVG(SELECT P.T.nurnshowingO FROM partition P)
FROM      Theaters T
GROUP BY  low: T.ticketPrice < 5, high: rr.ticketPrice >= 5
```

The GROUP BY clause now creates just two partitions called *low* and *high*. Each theater object T is placed in one of these partitions based on its ticket price. In the SELECT clause, lo'wand *high* are boolean variables, exactly one of which is true in any given output tuple; partition is instantiated to the corresponding partition of theater objects. In our exarnple, we get two result tuples. ()ne of thern has *low* equal to true and *avgNum* equal to the average nurnber of rnovies shown at theaters \vith a low ticket price. The second tuple has *high* equal to true and *avgNum* equal to the average nurnber of Inovies shown at theaters with a high ticket price.

The next query illustrates OQL support for queries that return collections other than set and rnultiset:

```
(SELECT   rr.theaterNarne
FROM      Theaters 'T
ORDER BY  T.ticketPrice DESC) [0:4]
```

The ORDER BY clause makes the result a list of theater naInes ordered by ticket price. The clcrnents of a list can be referred to by position, starting \vith position 0. Therefore, the expression [0:4] extracts a list containing the names of the five theaters \vith the highest ticket prices.

OQL also supports DISTINCT, HAVING, explicit nesting of subqueries, view definitions, and other SQL features.

23.10 COMPARING RDBMS, OODBMS, AND ORDBMS

Now that we have covered the lnain object-oriented DBMS extensions, it is tirne to consider the two lnain variants of object-databases, OODBMSs and ORDBJVISs, and cornpare thern with RDBMSs. Although we presented the concepts underlying object-databases, we still need to define the tenns OODBMS and ORDBMS.

An ORDBMS is a relational DBl\1S with the extensions discussed in this chapter. (Not all ORDBMS systerlls support all the extensions in the general forrn that we have discussed theIn, but our concern in this section is the paradigrll itself rather than specific systenls.) An OODBMS is a progranlrning language with a type system that supports the features discussed in this chapter and allows any data object to be persistent; that is, to survive across different prograrn executions. Many current systerns conform to neither definition entirely but are lIluch closer to one or the other and can be classified accordingly.

23.10.1 RDBMS versus ORDBMS

COlllparing an RDBMS with an OI{DBMS is straightforward. An R,DBl\1S does not support the extensions discussed in this chapter. rrhe resulting sirnplicity of the data rnodel rnakes it easier to optirnize queries for efficient execution, for example. A relational systern is also easier to use because there are fewer features to master. ()n the other hand, it is less versatile than an ()HJ)BiviS.

23.10.2 OODBMS versus ORDBMS: Similarities

OODBMSs and ()HI)BJVISs both support user-defined ADTs, structured types, ol)ject identity and reference types, and inheritance. Both support a query language for rnanipulating collection types. ORDBMSs support an extended fonn of SQL, and 001)131\18s support ODL/OQL. The similarities are by no rneans accidental ORDBMSs consciously try to add OODBMS features to an RI)B1\118, and OODBMSs in turn have developed query languages based on

relational query languages. Both OODBMSs and ORDBMSs provide DBMS functionality such as concurrency control and recovery.

23.10.3 OODBMS versus ORDBMS: Differences

The fundaulental difference is really a philosophy that is carried all the way through: OODBMSs try to add DBMS functionality to a progranllning language, whereas ORDBMSs try to add richer data types to a relational DBlVIS. Although the two kinds of object-databases are converging in terrns of functionality, this difference in their underlying philosophy (and for most systeIns, their irnplementation approach) has iInportant consequences in terIllS of the issues emphasized in the design of these DBJVISs and the efficiency with which various features are supported, as the following comparison indicates:

- OODBMSs airn to achieve seamless integration with a programrning language such as C++, Java, or Smalltalk. Such integration is not an important goal for an ORDBMS. SQL:1999, like SQL-92, allows us to embed SQL commands in a host language, but the interface is very evident to the SQL programer. (SQL:1999 also provides extended prograrnming language constructs of its own, as we saw in Chapter 6.)

- An OODBMS is aimed at applications where an object-centric viewpoint is appropriate; that is, typical user sessions consist of retrieving a few objects and working on theHl for long periods, with related objects (e.g., objects referenced by the original objects) fetched occasionally. Objects rnay be extrelnely large and rnay have to be fetched in pieces; therefore, attention Inust be paid to buffering parts of objects. It is expected that rnost applications can cache the objects they require in rnemory, once the objects are retrieved froIn disk. rrherefore, considerable attention is paid to rnaking references to ill-lnernory objects efficient. Transactions are likely to be of very long duration and holding locks until the end of a transaction Inay lead to poor perfonnance; therefore, alternatives to Two-Phase Locking HUlst be used.

  An ORDBMS is optirnized for applications in which large data collections are the focus, even though objects rnay have rich structure and be fairly large. It is expected that applications will retrieve data frorn disk extensively and optirnizing disk access is still the rnain concern for efficient execution. Transactions are assurned to be relatively short and traditional R.DB1VIS techniques are typically used for concurrency control and recovery.

- The query facilities of OQL are not supported efficiently in rnost OODBMSs, whereas the query facilities are the centerpiece of an ()HI)Bl\1S. To scnne extent, this situation is the result of different concentrations of effort in the clevelopment of these systerns. '1'0 a significant extent, it is also a

consequence of the systerlls' being optirnized. for very different kinds of applications.

23.11 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Consider the extended Dinky exarnple frorn Section 23.1. Explain how it lllotivates the need for each of the following object-database features: *'User-defined struct'Ured types, abstract data types (AD Ts), inheritance,* and *object identity.* (Section 23.1)
- What are *structured data types?* What are *collection types,* in particular? Discuss the extent to which these concepts are supported in SQL:1999. What irnportant type constructors are lllissing? What are the limitations on the ROWand ARRAY constructors? (Section 23.2)
- What kinds of operations should be provided for each of the structured data types? To what extent is such support included in SQL:1999? (Section 23.3)
- What is an *abstract data type?* How are nlethods of an abstract data type defined in an external programnling language? (Section 23.4)
- Explain *inheritance* and how new types (called *subtypes)* extend existing types (called *supertypes).* What are *rnethod overloading* and *late b'inding?* What is a *collect'ion hierarchy*? Contrast this with inheritance in prograrn-Ining languages. (Section 23.5)
- How is an *object identifier (aid)* different froln a record id in a relational DElVIS? How is it different froIn a URL? What is a *reference type?* Define *deep* and *shallow* equalit,y and illustrate thern through an exarnple. (Section 23.6)
- 'The rnultitude of data types in an ORDBMS allows us to design a rnore natural and efficient databa"se schema but introduces some new design choices. I)iscuss ORDBMS database design issues and illustrate your discussion using an exalnple application. (Section 23.7)
- Irnplernenting an ORDBMS brings new challenges. The systcrll rnust store large ADTs and structured types that rnight be very large. Efficient and extensible index rnechanisrns lllUSt be provided. Examples of new functionality include *user-defined aggregation functions* (we can define new aggregation functions for our AI)Ts) and *method security* (the systcrIl has to prevent user-defined rnethods fronl cornprornising the security of the DBMS). ExarIlples of new techniques to increase perfonnance include

method caching and *pointer swizzling*. The optirnizer must know about the new functionality and use it appropriately. Illustrate each of these challenges through an exarnple. (Section 23.8)

- Compare OODBIvISs with ORDBMSs. In particular, cornpare OQL and SQL:1999 and discuss the underlying data rnode!. (Sections 23.9 and 23.10)

EXERCISES

Exercise 23.1 Briefly answer the following questions:

1. What are the new kinds of data types supported in object-database systcrIls? Give an exarnple of each and discuss how the exanlple situation would be handled if only an RDBMS were available.
2. What rIlust a user do to define a new ADT?
3. Allowing users to define rIlethods can lead to efficiency gains. Give an exarnple.
4. What is late binding of nlCthods? Give an exarnple of inheritance that illustrates the need for dynamic binding.
5. What are collection hierarchies? Give an exalIlple that illustrates how collection hierarchies facilitate querying.
6. Discuss how a DBMS exploits encapsulation in ilnplernenting support for ADTs.
7. Give an exarnple illustrating the nesting and unnesting operations.
8. Describe two objects that are deep equal but not shallow equal or explain why this is not possible.
9. Describe two objects that are shallow equal but not deep equal or explain why this is not possible.
10. COlnpare RDBNISs with ORDBMSs. Describe an application scenario for which you would choose an RDBMS and explain why. Silnilarly, describe an application scenario for which you would choose an ORDBMS and explain why.

Exercise 23.2 Consider the Dinky schclna shown in Figure 23.1 and all related lncthocls defined in the chapter. Write the following queries in SQL:1999:

1. How luany filrns were shown at theater *tno* = 5 between January 1 and February 1 of 2002'1
2. What is the lowest budget for a filnl with at least two stars?
3. Consider theaters at which a fihu directed by Steven Spielberg started showing on January 1, 2002. For each such theater, print the narnes of all countries within a 100-ruile radius. (You can use the *overlap* and *radius* rnethods illustrated in Figure 2:3.2.)

Exercise 23.3 In a cornpany database, you need to store inforrnation about clnployees, departrnents, and children of erIlployees. For each ernployec, identified by *ssn*, you rnust record *years* (the number of years that the ernployee has worked for the cornpany), *phone*, and *photo* inforrnation. There are two subclasses of ernployees: contract and regular. Salary is coruputed by invoking a rnethod that takes *years* as a parameter; this rnethod has a different

irnplernentation for each subclass. Further, for each regular employee, you IIIUSt record the name and age of every child. The rnost conunon queries involving children are sirnilar to "Find the average age of Bob's children' and "Print the narnes of all of Bob's children."

A photo is a large image object and call be stored in one of several irnage fonnats (e.g., gif, jpeg). You want to define a *display* rnethod for image objects; display IIlust be defined differently for each irnage fonnat. 'For each department, identified by *dno*, you rnust record *dname*, *budget*, and *workers* infonnation. *Workers* is the set of crllployees who work in a given department. Typical querie.s involving workers include, "Find the average salary of all workers (across all departments)."

1. Using extended SQL, design an ORDBIVIS scherna for the cornpany database. Show all type definitions, including rnethod definitions.
2. If you have to store this infonnation in an RDBMS, what is the best possible design?
3. Cornpare the ORDBMS and RDBMS designs.
4. If you are told that a COlllInon request is to display the irnages of all employees in a given departruent, how would you use this inforulation for physical database design?
5. If you are told that an ernployee's ilnage rnust be displayed whenever any information about the employee is retrieved, would this affect your scherna design?
6. If you are told that a COU1Inon query is to find all erTlployees who look sirnilar to a given image and given code that lets you create an index over all irnages to support retrieval of sinlilar images, what would you do to utilize this code in an OR.DBMS?

Exercise 23.4 ORDBMSs need to support efficient access over collection hierarchies. Consider the collection hierarchy of Theaters and Theater_cafes presented in the Dinky exanlple. In your role as a DBMS illlplernentor (not a DBA), you rnust evaluate three storage alternatives for these tuples:

- All tuples for all kinds of theaters are stored together all disk in an arbitrary order.
- All tuples for all kinds of theaters are stored together on disk, with the tuples that are frOIll TheateLcafes stored directly after the last of the non-cafe tuples.
- T'uples froIll Theater_cafes are stored separately froIll the rest of the (non-cafe) theater tuples.

1. F'or each storage option, describe a rnechanisrn for distinguishing plain theater tuples frorn Theater_cafe tuples.
2. For each storage option, describe hmv to handle the insertion of a new non-cafe tuple.
3. Which storage option is 1110St efficient for queries over all theaters? Over just TheateLcafes? In terrns of the nurnber of 1/Os, how rnuch rnore efficient is the best technique for each type of query cornpared to the other two techniques?

Exercise 23.5 Different ORDBMSs use different techniques for building indexes to evaluate queries over collection hierarchies. For our Dinky example, to index theaters by name there are two COIIIIIIon options:

- Build one 13+ tree index over Theaters.*name* and another 13+ tree index over Theater_cafes.*narne*.
- Build one B+ tree index over the union of Theaters.*name* and Theater_cafes.*name*.

1. Describe how to efficiently evaluate the following query using each indexing option (this query is over aU kinds of theater tuples):

   SELECT * FROM Theaters T WHERE T.narne= 'Majestic'

   Give an estimate of the nurnber of l/Os required in the two different scenarios, assurning there are 1 rnillion standard theaters and 1000 theater-cafes. Which option is Inore efficient?

2. Perforrn the sallIe analysis over the following query:

   SELECT * FROM Theater-cafes 'I' WHERE T.nalne = 'Majestic'

3. For clustered indexes, does the choice of indexing technique interact with the choice of storage options? For unclustered indexes?

Exercise 23.6 Consider the following query:

```
SELECT thurnbnail(Lirnage)
FROM   lInages I
```

Given that the 1.*image* colurnn 111ay contain duplicate values, describe how to use hashing to avoid conlputing the *thum,bnail* function rnore than once per distinct value in processing this query.

Exercise 23.7 You are given a two-dimensional, n x n array of objects. Assume that you can fit 100 objects on a disk page. Describe a way to layout (chunk) the array onto pages so that retrievals of square m x m subregions of the array are efficient. (Different queries request subregions of different sizes, i.e., different m values, and your arrangement of the array onto pages should provide good perforrnance, on average, for all such queries.)

Exercise 23.8 An ORDBMS optiruizer is given a single-table query with n expensive selection conditions, $\sigma_{n(\ldots(\sigma_1(T)))}$. For each condition σ_i, the optirnizer can estinlate the cost c_i of evaluating the condition on a tuple and the reduction factor of the condition r_i. Assurne that there are t tuples in T.

1. How many tuples appear in the output of this query?
2. Assurning that the query is evaluated as shown (without reordering selections), what is the total cost of the query? Be sure to include the cost of scanning the table and applying the selections.
3. In Section 23.8.2, it was asserted that the optiruizer should reorder selections so that they are applied to the table ill order of increasing rank, where $\text{rank}_i = (r_i - 1)/c_i$. Prove that this assertion is optirual. That is, show that no other ordering could result in a query of lower cost. (Hint: It may be easiest to consider the special case where $n = 2$ first and generalize from there.)

Exercise 23.9 ORDBIVlSs support references as a data type. It is often clailnecl that using references instead of key–foreign key relationships will give rnuch higher perfonnance for joins. 'This question asks you to explore this issue.

- Consider the following SQL:1999 DDL which only uses straight relational constructs:

  ```
  CREATE TABLE R(rkey integer, rdata text);
  CREATE TABLE S(skey integer, rfkey integer);
  ```

Assurne that we have the following straightforward join query:

```
SELECT  S.skey, H..relata
FROM    S, R
WHERE   S.rfkey = R.rkey
```

- Now consider the following SQL:1999 ORDBMS schelna:

```
CREATE TYPE r_t AS ROW(rkey integer, rdata text);
CREATE TABLE R OF r_t REF is SYSTEM GENERATED;
CREATE TABLE S (skey integer, r REF(r_t) SCOPE R);
```

Assurne we have the following query:

```
SELECT  S.skey, S.r.rkey
FROM    S
```

What algorithrll would you suggest to evaluate the pointer join in the ORDBMS scherna? How do you think it will perform versus a relational join on the previous scherna?

Exercise 23.1.0 Many object-relational systerns support set-valued attributes using some variant of the setof constructor. For eXaInple, assurning we have a type **person_t**, we could have created the table Filrns in the Dinky Schema in Figure 23.1 as follows:

CREATE TABLE *Films(filrnno* integer, *title* text, *stars* setof Person);

1. Describe two ways of irnpleIIlenting set-valued attributes. One way requires variable-length records, even if the set elements are all fixed-length.
2. Discuss the irnpact of the two strategies on optimizing queries with set-valued attributes.
3. Suppose you would like to create an index on the column stars in order to look up filrns by the narne of the star that has starred in the fillll. For both irnplenlentation strategies, discuss alternative index structures that could help speed up this query.
4. What types of statistics should the query optirnizer rnaintain for set-valued attributes? How do we obtain these statistics'?

BIBLIOGRAPHIC NOTES

A nurnber of the object-oriented features described here are based in part on fairly old ideas in the prograrnrning languages cornrnunity. [42] provides a good overview of these ideas in a database context. Stonebraker's book [719J describes the vision of ORDBMSs ernbodied by his company's early product, Illustra (now a product of Inforrnix). Current connnercial DBMSs with object-relational support include Infonnix Universal Server, IBM D13/2 CS V2, and UniSQL. An new version of Oracle is scheduled to include ORDBMS features as well.

Many of the idcas in current object-relational systerlls carne out of a few prototypes built in the 1980s, especially POSTGRES [723], Starburst [351], and 02 [218].

The iclea of an object-oriented database was first articulated in [197], \vhich described the GernStone prototype system. Other prototypes includeDASDBS [657], EXODUS [130], nus [273], Ol:>jectStore [463], ODE, [18] ORION [432], SHORE [129], and THOR [482]. 02 is actually an early example of a systenl that was beginning to rnerge the thcrnes of ORDBrvISs

and OODBMSs—it could fit in this list as well. [41] lists a collectioll of features that are generally considered to belong in an OODBMS. Current cornrnercially available OODBMSs include GelllStone, Itasca, 02, Objectivity, ObjectStore, Ontos, Poet, and Versant. [431] cornpares OODBIvISs and RDBMSs.

Database support for ADTs was first explored in the INGRES and POSTGRES projects at U.C. Berkeley. The basic ideas are described in [716], including Inechanisllls for query processing and optilnization with ADTs as well as extensible indexing. Support for ADTs was also investigated in the Dannstadt database systern, [480]. Using the POSTGRES index extensibility correctly required intimate knowledge of DBTvIS-internal transaction ruechanisllls. Generalized search trees were proposed to solve this problern; they are described in (376], with concurrency and ARIES-based recovery details presented in [447]. [672] proposes that users lnust be allowed to define operators over ADT objects and properties of these operators that can be utilized for query optiInization, rather than just a collection of lnethods.

Array chunking is described in (653]. Techniques for luethod caching and optimizing queries with expensive lnethods are presented in [37:3, 165]. Client-side data caching in a client-server 00D131\I1S is studied in [283]. Clustering of objects on disk is studied in [741]. Work on nested relations was an early precursor of recent research on complex objects in OODBMSs and ORD13IvISs. One of the first nested relation proposals is (504]. MVDs play an inlportant role in reasoning about reduncancy in nested relations; see, for exalnple, [579]. Storage structures for nested relations were studied in (215].

Fonnal rnodels and query languages for object-oriented databases have been widely studied; papers include [4, 56, 75, 125, 391, 392, 428, 578, 724]. [427] proposes SQL extensions for querying object-oriented databases. An early and elegant extension of SQL with path expressions and inheritance was developed in GEM [791]. There has been ITluch interest in cornbining deductive and object-oriented features. Papers in this area include (44, 288, 495, 556, 706, 793]. See [3] for a thorough textbook discussion of fonnal aspects of object-orientation and query languages.

[433, 435, 721, 796] include papers on DBMSs that would now be tenned object-relational and/or object-oriented. [794] contains a detailed overview of scherna and database evolution in object-oriented database systenls. A thorough presentation of SQL:1999 can be found in [525), and advanced features, including the object extensions, are covered in [523]. A short survey of new SQL:1999 features can be found in [2:37]. The incorporation of several SQL:1999 features into I131v1 D132 is described in [128J. OQL is described in [141]. It is based to a large extent on the 02 query language, which is described, together with other aspects of 02, in the collection of papers [55].

24

DEDUCTIVE DATABASES

- What is the nlotivation for extending SQL with recursive queries?
- What important properties must recursive programs satisfy to be practical?
- What are least lnodels and least fixpoints and how do they provide a theoretical foundation for recursive queries?
- What cOlnplications are introduced by negation and aggregate operations? How are they addressed?
- What are the challenges in efficient evaluation of recursive queries?
- Key concepts: Datalog, deductive databases, recursion, rules, infeI'ences, safety, range-restriction; least model, declarative semantics; least fixpoint, operational semantics, fixpoint operator; negation, stratified program.s; aggregate operators, rnultiset generation, grouping; efficient evaluation, avoiding repeated inferences, Seminaive fixpoint evaluation; pushing query selections, lVlagic Sets rewriting

For 'Is' and 'Is-Not' though with Rule and Line,
And 'Up-and-Down' by Logic I define,
Of all that one should care to fathorn, I
Was never deep in anything but-------\Vine.

Rubaiyat of Omar !(hayyarn, Translated by Edward Fitzgerald

Relational database rnanagernent systenls have been enonnously successful for C),chninistrative da,ta processing. In recent years, however, as people have tried to

use database systerus in increasingly cornplex applications, some irnportant linlitations of these systellls have been exposed. For sonle applications, the query language and constraint definition capabilities have been found inadequate. As an exarnple, sonle cornpanies ulaintain a huge parts inventory database and frequently want to ask questions such as, "Are we running Iowan any parts needed to build a ZX600 sports car?" or "What is the total cornponent and assernbly cost to build a ZX600 at today's part prices?" These queries cannot be expressed in SQL-92.

We begin this chapter by discussing queries that cannot be expressed in relational algebra or SQL and present a rnore powerful relational language called *Datalog.* Queries and views in SQL can be understood as **if–then** rules: "**If** some tuples exist in tables mentioned in the FROM clause that satisfy the conditions listed in the WHERE clause, **then** the tuple described in the SELECT clause is included in the answer." Datalog definitions retain this if-then reading, with the significant new feature that definitions can be *recursive,* that is, a table can be defined in terms of itself. The SQL:1999 standard, the successor to the SQL-92 standard, requires support for recursive queries, and a large subset S0111e systerlls, notably IBM's DB2 DBMS, already support thelu.

Evaluating Datalog queries poses some additional challenges, beyond those encountered in evaluating relational algebra queries, and we discuss sonle iUlportant ilnplernentation and optimization techniques developed to address these challenges. Interestingly, some of these techniques have been found to irnprove perforrnance of even nonrecursive SQL queries and have therefore been implernented in several current relational DBMS products.

In Section 24.1, we introduce recursive queries and Datalog notation through an exaruple. We present the theoretical foundations for recursive queries, least fixpoints and least rnodels, in Section 24.2. We discuss queries that involve the use of negation or set-difference in Section 24.3. Finally, we consider techniques for evaluating recursive queries efficiently in Section 24.5.

24.1 INTR()DUCTION TO RECURSIVE QUERIES

We begin with a sinlple example that illustrates the li111its of SQL-92 queries and the power of recursive definitions. Let Assernbly be a relation \vith three fields *part*, *subpart*, and *qty.* An example instance of Assernbly is shown in Figure 24.1. Each tuple in Assernbly indicates IH}w Inany copies of a particular subpart are COlltained in a given part. The first tuple indicates, for example, that a trike contains three wheels. The Assclnbly relation can be visuaJized as a tree, as sho\vn in Figure 24.2. A tuple is shown as an edge going frorn the part to the subpart, with the *qty* value as the edge label.

part	*subpart*	*qty*
trike	\vheel	3
trike	fralne	1
frarne	seat	1
franle	pedal	1
wheel	spoke	2
\vheel	tire	1
tire	run	1
tire	tube	1

Figure 24.1 An Instance of Assembly

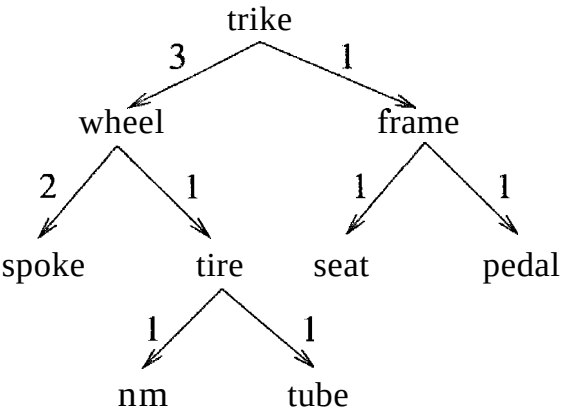


Figure 24.2 Assembly Instance Seen as a Tree

A natural question to ask is, "What are the cornponents of a trike?" Rather surprisingly, this query is inlpossible to write in SQL-92. Of course, if we look at a given instance of the Assernbly relation, we can write a 'query' that takes the union of the parts that are used in a trike. But such a query is not interesting---we want a query that identifies all components of a trike for *any* instance of Assembly, and such a query cannot be written in relational algebra or in SQL-92. Intuitively, the problem is that we are forced to join the Asselnbly relation with itself to recognize that *trike* contains *spoke* and *tire,* that is, to go one level down the Assenlbly tree. For each additional level, we need an additional join; two joins are needed to recognize that *trike* contains *rim,* which is a subpart of *tire.* Thus, the ntullber of joins needed to identify all subparts of *trike* depends on the height of the Assen1bly tree, that is, on the given instance of the Assembly relation. No relational algebra query works for all instances; given any query, we can construct an instance whose height is greater than the nurnber of joins in the query.

24.1.1 Datalog

We now define a relation called Cornponents that identifies the cOlnponents of every part. Consider the following program, or collection of rules:

```
Components(Part, SUbpart) "- Assembly(Part, SUbpart, Qty) "
Components(Part, Subpart) .- Assembly(Part, Part2, Qty),
                             Components(Part2, Subpart)"
```

These are rules in Datalog, a relational query language inspired by Prolog, the well-known logic progranuning language; indeed, the notation follows Prolog. The first rule should be read as follo\vs:

For all values of Part, Subpart, and Qty,

if there is a tuple (Part, Subpart, Qty⟩ in Assclnbly,
then there IJIUSt be a tuple ⟨Part, Subpart) in (;olnponents.

The second rule should be read as follows:

For all values of Part, Part2, Subpart, and Qty,
if there is a tuple (Part, Part2, Qty⟩ in Assernbly and
a tuple ⟨Part2, Subpart) in Components,
then there HUlst be a tuple ⟨Part, Subpart) in C()lnponents.

The part to the right of the :- sYlnbol is called the body of the rule, and the part to the left is called the head of the rule. The syrnbol :- denotes logical irnplication; if the tuples lIlentioned in the body exist in the database, it is irnplied that the tuple rnentioned in the head of the rule rnust also be in the database. (Note that the body could be ernpty; in this case, the tuple rnentioned in the head of the rule rnust be included in the database.) 1'herefore, if we are given a set of Assenlbly and Cornponents tuples, each rule can be used to infer, or deduce, sorne new tuples that belong in COlnponents. This is why database systerns that support Datalog rules are often called deductive database systems.

By assigning constants to the variables that appear in a rule, we can infer a specific Coruponents tuple. For example, by setting Part=*trike,* Subpart=*wheel,* and Qty=*3*, we can infer that ⟨*trike*, *wheel)* is in eoulponents. Each rule is really a *ternplate* for Inaking inferences: An inference is the use of a rule to generate a new tuple (for the relation in the head of the rule) by substituting constants for varia,bles in such a way that every tuple in the rule body (after the substitution) is in the corresponding relation instance.

By considering each tuple in Asselnbly in turn, the first rule allows us to infer that the set of tuples obtained by taking the projection of Assernbly onto its first two fields is in CCHnponents.

The seco11d rule then allo\vs us to cOlnbine previously discovered Cornponents tuples with Assernbly tuples to infer new Cornponents tuples. We can apply the second rule by considering the cross-product of Assernbly and (the current instance of) Cornponents and assigning values to the variables in the rule for each row of the cross-product, one row at a time. ()bserve how the repeated use of the varial)le Part2 prevents certain rows of the cross-product fronl contributing any new tuples; in effect, it specifies an equality join condition on AssenIbly and Cornpouents. The tuples obtained by one application of this rule are shown in Figure 24.3. (In addition, COlnponents contains the tuples obtained by applying the first rule; these are not shown.)

part	*subpart*
trike	spoke
trike	tire
trike	seat
trike	pedal
wheel	rhn
wheel	tube

Figure 24.3 Components Tuples Obtained by Applying the Second Rule Once

part	*subpart*
trike	spoke
trike	tire
trike	seat
trike	pedal
wheel	run
whe81-	tube
trike	rirn
trike	tube

Figure 24.4 Components Tuples Obtained by Applying the Second Rule Twice

The tuples obtained by a second application of this rule are shown in Figure 24.4. Note that each tuple shown in Figure 24.3 is reinferred. Only the last two tuples are new.

Applying the second rule a third time does not generate additional tuples. rrhe set of Components tuples shown in Figure 24.4 includes all the tuples that can be inferred using the two Datalog rules defining Cornponents and the given instance of Assembly. rrhe components of a trike can now be obtained by selecting all Cornponents tuples with the value *trike* in the first field.

Each application of a Datalog rule can be understood in ternlS of relational algebra. The first rule in our exarnple program simply applies projection to the Assernbly relation and adds the resulting tuples to the Cornponents relation, which is initially ernpty. The second rule joins Assernbly with COlllponents and then does a projection. The result of each rule application is cornbined with the existing set of Cornponents tuples using union.

The only Datalog operation that goes beyond relational algebra is the *repeated* application of the rules defining CCHnponents until no new tuples are generated. This repeated application of a set of rules is called the *fixpoint* operation, and we develop this idea further in the next section.

We conclude this section by rewriting the Datalog definition of Cornponents using SQL:1999 syntax:

```
WITH RECURSIVE Cornponents(Part, Subpart) AS
        (SELECT A1.Part, A1.Subpart FROM Assernbly A1)
        UNION
        (SELECT A2.Part, C1.Subpart
        FROM    Assernbly A2, Cornponents C1
```

```
        WHERE  A2.Subpart = C1.Part)

SELECT * FROM COlllponents C2
```

The WITH clause introduces a relation that is part of a query definition; this relation is shnilar to a view, but the scope of a relation introduced using WITH is local to the query definition. The RECURSIVE key\vord signals that the table (in our example, Cornponents) is recursively defined. The structure of the definition closely parallels the Datalog rules. Incidentally, if we wanted to find the cornponents of a particular part, for exanlple, *tTikc,* we can sirnply replace the last line with the following:

```
SELECT * FROM Cornponents C2
WHERE  C2.Part = 'trike'
```

24.2 THEORETICAL FOUNDATIONS

We classify the relations in a Datalog prograln as either output relations or input relations. Output relations are defined by rules (e.g., COluponents), and input relations have a set of tuples explicitly listed (e.g., Assembly). Given instances of the input relations, we Inust compute instances for the output relations. The meaning of a Datalog prograrIl is usually defined in two different ways, both of which essentially describe the relation instances for the output relations. Technically, a query is a selection over one of the output relations (e.g., all Components tuples C with C.*paTt = tTike).* However, the lueaning of a query is clear once we understand how relation instances are associated with the output relations in a Datalog progranl.

rrhe first approach to defining the sernantics of a Datalog prograllll, called the *least model semantics*, gives users a way to understand the prograrn without thinking about how the prograrn is to be executed. That is, the sernanties is *declarative*, like the sernantics of relational calculus, and not *operational* like relational algebra sClnantics. This is irnportant becC111se recursive rules lnake it difficult to understand a· prograIll in tcrrns of an evaluation strategy.

The second approach, called the *least fixpoint 8crnantic8,* gives a conceptual evaluation strategy to COlnpute the desired relation insta.nces. This serves as the basis for recursive query evaluation in a DBMS. More efficient evaluation strategies are used in an actual iInplernentation, but their correctness is sho\vI1 by demonstrating their equivalence to the least fixpoint approach. rrhe fixpoint sClnantics is thus operational and. plays a role analogous to that of relational algebra sernalltics for nonrecursive queries.

24.2.1 Least Model Semantics

We want users to be able to understand a Datalog progTarn by understanding each rule independent of other rules, with the lneaning: *If the body is true, the head is also true.* This intuitive reading of a rule suggests that, given certain relation instances for the relation naines that appear in the body of a rule, the relation instance for the relation rnentioned in the head of the rule 111USt contain a certain set of tuples. If a relation Harne R appears in the heads of several rules, the relation instance for R IIIUSt satisfy the intuitive reading of all these rules. However, we do not want tuples to be included in the instance for R unless they are necessary to satisfy one of the rules defining R. That is, we want to cornpute only tuples for R that are supported by SaIne rule for R.

To lnake these ideas precise, we need to introduce the concepts of rnodels and least models. A model is a collection of relation instances, one instance for each relation in the prograrn, that satisfies the following condition. For every rule in the prograrll, whenever we replace each variable in the rule by a corresponding constant, the following holds:

> *If* every tuple in the body (obtained by our replaceUlent of variables with constants) is in the corresponding relation instance,
>
> *Then* the tuple generated for the head (by the assignrnent of constants to variables that appear in the head) is also in the corresponding relation instance.

Observe that the instances for the input relations are given, and the definition of a rnodel essentially restricts the instances for the output relations.

Consider the rule

```
Components(Part, Subpart) :- Assembly(Part, Part2, Qty),
                             Components(Part2, Subpart).
```

Suppose we replace the variable Part by the constant *wheel,* Part2 by *tire,* Qty by 1, and Subpart by *rim:*

```
Components(wheel, rim) :- Assembly(wheel, tire, 1),
                          Components(tire, rim).
```

Let A be an instance of Assernbly and C be an instance of COlnpouents. If A contains the tuple $\langle wheel, tire, 1\rangle$ and C contains the tuple $\langle tire, rim\rangle$, then C rrulst also contain the tuple $\langle wheel, rim\rangle$ for the pair of instances A and C

to be a rnodel. ()f course, the instances A and C rnust satisfy the inclusion requirenlent just illustrated for *every* assignruent of constants to the variables in the rule: If the tuples in the rule body are in.A. and C, the tuple in the head Inus1; be in C.

As an exarnple, the instances of Asscrnbly shown in Figure 24.1 and Cornponents shown in F'igure 24.4 together fornl a rnodel for the Conlponcnts prograll1.

C;iven the instance of Assernbly shown in Figure 24.1, there is no justification for including the tuple ⟨*spoke*, *pedal)* to the COlnponents instance. Indeed, if we add this tuple to the cornponents instance in Figure 24.4, we no longer have a lllodel for our program, as the following instance of the recursive rule derllonstrates, since *(wheel, pedal)* is not in the Cornponents instance:

```
Components(wheel, pedal) :- Assembly(wheel, spoke, 2),
                            Components(spoke, pedal).
```

However, by also adding the tuple ⟨*wheel*, *pedal)* to the Cornponents instance, we obtain another rnodel of the Components prograrll. Intuitively, this is unsatisfactory since there is no justification for adding the tuple ⟨*spoke*, *pedal)* in the first place, given the tuples in the Assembly instance and the rules in the prograln.

We address this problern by using the concept of a least rllodel. A least model of a prograrn is a rnodel M such that for every other model M2 of the sarne prograrnl, for each relation R in the program, the instance for R in ℕ is contained in the instance of R in 1\12. The 1nodel forIned by the instances of Assernbly and COlnponents shown in Figures 24.1 and 24.4 is the least rHodel for the CC)lnponents prograllll with the given Assernbly instance.

24.2.2 The Fixpoint Operator

A fixpoint of a function *f* is a value *v* such that the function applied to the value returns the saIIle value, that is, *f(v)* = *v*. Consider a function applied to a set of values that also returns a set of values. For example, we carl define *double* to be a function tllat Illuitiplies every element of the input set by two and *double+* tobe *double* ∪ *identity*. Thus, *double*({1,2,5}) = {2,4,lO}, and *double+*({1,2,5}) = {1,2,4.,5,lO}.The set of all even integers which happens to be an infinite set-is a fixpoint of the function *double+*. Another fixpoint of the function *double+* is the set of all integers. The first fixpoint (the set of all even integers) is *smaller* than the second fixpoint (the set of all integers) because it is contained in the latter.

The least fixpoint of a function is the fixpoint that is slnaller than every other fixpoint of that function. In general, it is not guaranteed that a function has a least fixpoint. For example, there lnay be two fixpoints, neither of \vhich is 81na11er than the other. (Does *double* have a least fixpoint? What is it?)

No\V let us turn to functions over sets of tuples, in particular, functions defined using relational algebra expressions. The Cornponents relation can be defined by an equation of the fonn

$$Components = \pi_{1,5}(Assembly \bowtie_{2=1} Components) \cup \pi_{1,2}(Assernbly)$$

1'his equation has the forn1

$$Cornponents = f(Cornponents, Assembly)$$

where the function f is defined using a relational aJgebra expression. For a given instance of the input relation Assernbly, this can be sirnplified to

$$Components = f(Co1nponents)$$

The least fixpoint of f is an instance of Cornponents that satisfies this equation. Clearly the projection of the first two fields of the tuples in the given instance of the input relation Assernbly rnust be included in the (instance that is the) least fixpoint of Cornponents. In addition, any tuple obtained by joining Components with Assernbly and projecting the appropriate fields IllUst also be in Components.

A little thought shows that the instance of Components that is the least fixpoint of f can be ccnnputed using repeated applications of the Datalog rules sho\vn in the previous section. Indeed, applying the two Datalog rules is identical to evaluating the relational expression used in defining COlnponcnts. If an application generates Cornponents tuples that are not in the current instance of the Cornponents relation, the current instance cannot be the fixpoint. Therefore, we add the new tuples to Cornponents and evaluate the relational expression (equivalently, the two Datalog rules) again. T'his process is repeated until every tuple generated is already in the current instance of Cornponents. When applying the rules to i,he currerlt set of tuples does not produce any new tuples, we have reached a fixpoint. If CC)lnponents is initialized to the erupty set of tuples. intuitively we infer only tuples that (1,1'e necessary by the definition of a fixpoint, and the fixpoint cornputed is the least fixpoint.

24.2.3 Safe Datalog Programs

Consider the followving p1'ograrn:

```
ComplexYarts(Part) :- Assembly(Part, Subpart, Qty), Qty > 2.
```

According to this rule, a cOlnplex part is defined to be any part that has Inore than tvo copies of anyone subpart. For each part lnentioned in the Asselnbly relatioIl, we can easily check whether it is a cOll1plex part. In contrast, consider the following prograrn:

```
PriceYarts(Part, Price) '-
        Assembly(Part, Subpart, Qty), Qty> 2.
```

This variation seeks to associate a price with each cornplex part. IIowever, the variable *Price* does not appear in the body of the rule. This Ineans that an infinite number of tuples must be included in any model of this progralll. To see this, suppose we replace the variable Part by the constant *trike,* SubPart by *wheel,* and Qty by 3. This gives us a version of the rule with the only remaining variable being *Price:*

```
PriceYarts(trike,Price) :- Assembly(trike, wheel, 3), 3 > 2.
```

Now, any assignment of a constant to *Price* gives us a tuple to be included in the output relation Price_.Parts. For example, replacing *Price* by 100 gives us the tuple Price_Parts(trike,lOO). If the least Inodel of a prograllI is not finite, for even one instance of its input relations, then we say the program is unsafe.

Database systems disallow unsafe programs by requiring that every variable in the head of a rule also appear in the body. Such prograins are said to be range-restricted, and every range-restricted Datalog prograin has a finite least model if the input relation instances are finite. In the rest of this chapter, we assume that prograrns are range-restricted.

24.2.4 Least Model =Least Fixpoint

Does a Datalog prograin always have a least rnodel? ()r is it possible that there are two rnodels, neither of which is contained in the other'? Sirnilarly, does every Datalog prograinl have a least fixpoint? VvThat is the relationship between the least rnodel and the least fixpoint of a Datalog prograin?

As we noted earlier, not every function has a least fixpoint. Fortunately, every function defined in terms of relational algebra expressions tl1at do not contain set-difference is ,guaranteed to have a least fixpoint, and the least fixpoint can be cornputed by repeatedly evaluating the functic)ll. This tells us that every l)atalog prograrn has a least fixpoint and that it can l)e cOlnputed by repeatedly applyi11g the rules of the l)rograinl on the given instances of the input relations.

Further, every Datalog program is guaranteed to have a least rnodel and the least rnodel is equal to the least fixpoint of the I>l'ograin. These results (whose

proofs we do not discuss) provide the basis for Datalog query processing. 'Users can understand a progTarn in terms of 'If the body is true, the head is also true,' thanks to the least l1lodel sClnantics. The DBMS can COlllpute the answer by repeatedly applying the prograrn rules, thanks to the least fixpoint sernantics and the fact that the least nlodel and the least fixpoint are identical.

24.3 RECURSIVE QUERIES WITH NEGATION

Unfortunately, once set-difference is allo\ved in the body of a rule, there r11ay be no least rnodel or least fixpoint for a program. Consider the following rules:

```
Big(Part):-     Assembly(Part, Subpart, Qty), Qty> 2,
                NOT Small(Part).
Small(Part) :-  Assembly(Part, Subpart, Qty), NOT Big(Part).
```

These two rules can be thought of as an attenlpt to divide parts (those that are mentioned in the first colulnn of the Asselubly table) into two classes, Big and Small. The first rule defines Big to be the set of parts that use at least three copies of some subpart and are not classified as small parts. The second rule defines Small as the set of parts not classified as big parts.

If we apply these rules to the instance of Assembly shown in Figure 24.1, *trike* is the only part that uses at least three copies of senne subpart. Should the tuple *(trike)* be in Big or SUlall? If we apply the first rule and then the second rule, this tuple is in Big. To apply the first rule, we consider the tuples in Asselubly, choose those with Qty > 2 (which is just *(trike)),* discard those in the current instance of Srnal1 (both Big and Small are initially elnpty), and add the tuples that are left to Big. 1'herefore, an application of the first rule adds *(trike)* to Big. Proceeding siInilarly, we can see that if the second rule is applied before the first, $\langle trike \rangle$ is added to Srnall instead of Big.

This program has two fixpoints, neither of 'which is srnaller than the other, as shown in Figure 24.5. (rhe first fixpoint has a Big tuple that does not appear in the second fixpoint; therefore, it is not smaller than the second fixpoint. The second fixpoint has a 81ma11 tuple that does not appear in the first fixpoint; therefore, it is D.ot sr11a11e1' than the first fixpoint. The order ill \vhich we apply the rules detennines \vhich fixpoint is cOlnputed; this situation is very unsatisfactory.\Ve want users to be able to understand their queries without thinking (1)out exactly ho\v the evaluation proceeds.

The root of the problerH is the use of NOT. When we apply the first rule, senne irlferences are disallowed because of the presence of tuples in 8rna11. Parts

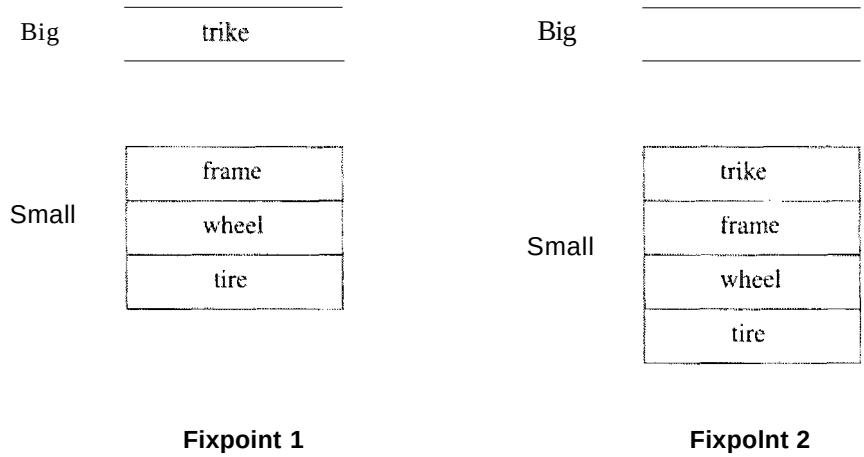


Figure 24.5 Two Fixpoints for the Big/Small Program

that satisfy the other conditions in the body of the rule are candidates for addition to Big; we remove the parts in 8rna11 frorn this set of candidates. Thus, sorne inferences that are possible if 8ruall is ernpty (as it is before the second rule is applied) are disallowed if SInall contains tuples (generated by applying the second rule before the first rule). Here is the difficulty: If NOT is used, the addition of tuples to a relation can *disallow* the inference of other tuples. Without NOT, this situation can never arise; the addition of tuples to a relation can *never* disallow the inference of other tuples.

Range-Restriction and Negation

If rules are allowed to contain NOT in the body, the definition of range-restriction rnust be extended ensure that all range-restricted prograrJlS are safe. If a relation appears in the body of a rule preceded by NOT, we call this a negated occurrence. Relation occurrences in the body that are not negated are called positive occurrences. A prograrn is range-restricted if every variable in the head of the rule appears in sorne positive relation occurrence in the body.

24.3.1 Stratification

A widely used solution to the problern caused by negation, or the use of NOT, is to irnpose certain syntactic restrictions on prograrlls. rrhese restrictions can be easily checked and programs that satisfy them have a natural meaning.

We say that a tableT depends on a table S if sorne rule with T in the head contains S, or (recursively) contains a predicate that depends on S, in the bod:y. A recursively defined predicate always depends on itself. For example, Big depends on Sruall (and on itself). Indeed, the tables Big and Srnall (l,re

nlutually recursive, that is, the definition of 'Big depends on SrnaU and vice versa. We say that a table T depends negatively on a table S if some rule with T in the head contains NOT S, or (recursively) contains a predicate that depends negatively on S, in the body.

Suppose we classify the tables in a prograrll into strata or layers as follows. The tables that do not depend on any other tables are in straturll 0. In our Big/S1nall exarnple, ASSCIIlbly is the only table in stratu1Il 0. Next, we identify tables in straturll 1; these are tables that depend only on tables in stratuln 0 or straturn 1 and depend negatively only on tables in straturn 0. Higher strata are sirnilarly defined: The tables in straturni are those that do not belong to lower strata, depend only on tables in stratuIll i or lower strata, and depend negatively only on tables in lower strata. A stratified program is one whose tables can be classified into strata according to the above algoritlull.

rrhe Big/Sruall prograIl is not stratified. Since Big and Snlall depend on each other, they 1nust be in the sarne straturn. Ho\vever, they depend negatively on each other, violating the requirc1Ilent that a table can depend negatively only on tables in lower strata. Consider the following variant of the Big/Srnall progra1Il, in which the first rule has been rnodified:

```
Big2(Part) :- Assembly(Part, Subpart, Qty), Qty> 2.
Smal12(Part) :- Assembly(Part, Subpart, Qty), NOT Big2(Part).
```

This prograrn is stratified. Slnall2 depends on Big2 but Big2 does not depend on 8111a1l2. Assernbly is in stratu111 0, Big is in straturn 1, and Srna1l2 is in straturn 2.

A stratified prograrn is evaluated stratu1n-by-straturn, starting with stratunl 0. 'To evaluate a straturn, we cornpllte the fixpoint of all rules defining tables in this straturn. When evaluating a straturn, any occurrence of NOT involves a table frorH a lower straturn, which has therefore been corupletely evaluated by now. The tuples in the negated table still disallow sorne inferences, but the effect is cornpletely deterrninistic, given the straturn-by-straturn evaluation. In the example, Big2 is C01uput8(1 before 81na1l2 because it is in a lower straturIl than 8rna112: $\langle trike \rangle$ is added to Big2. Next, 'when we cornpute 81na112, we recognize that $\langle trike \rangle$ is not in 8rna112 l)ecause it is already in Big2.

Incidentally, note that the stratified Big/Srnall prograrnl is not even recursive. If we replace Assernbl.y by Cornponents, we obtain a recursive, stratified prograrn: }\,sscrnbly is in straturn 0, Cornponents is in stratlull 1, Big2 is also in stratum 1, and 81na112 is in straturn 2.

Intuition behind Stratification

Consider the stratified version of the Big/Slnall prograrll. The rule defining Big2 forces us to add $\langle trike \rangle$ to Big2 and it is natural to assume that $\langle trike \rangle$ is the only tuple in Big2, because we have no supporting evidence for any other tuple being in Big2. The rninirnal fixpoint conrputecl by stratified fixpoint evaluation is consistent \vith this intuition. However, there is another rninhnal fixpoint: We can place every part in Big2 and rnake Sma1l2 be ernpty. While this assignrllent of tuples to relations seeIns unintuitive, it is nonetheless a rninimal fixpoint.

rrhe requirernent that prograrns be stratified gives us a natural order for evaluating rules. When the rules are evaluated in this order, the result is a unique fixpoint that is one of the minirnal fixpoints of the progrArll. The fixpoint C0111puted by the stratified fixpoint evaluation usually corresponds well to our intuitive reading of a stratified progrArll, even if the program has rnore than one rllininlal fixpoint.

For nonstratified Da.talog progranls, it is harder to identify a natural model frorn arnong the alternative rninirnal rnodels, especially when we consider that the Ineaning of a progrArll must be clear even to users who lack expertise in Dlathelnatical logic. Although considerable research has been done on identifying natural rnodels for nonstratified prograrns, practical irnplernentations of Datalog have concentrated on stratified prograrns.

Relational Algebra and Stratified Datalog

Every relational algebra query can be written as a range-restricted, stratified Datalog progra.rn. (Of course, not all Datalog progranls can be expressed in relational algebra; for exarnple, the Cornponents prograrn.) 'We sketch the translation frorn algebra to stratified Datalog by writing a Datalog progra.rn for each of the basic algebra operations, in terrns of two eXC1mple tables R and S, each with two fields:

Selection:	Result(Y) :- Il(X,Y), X=c.
Projection:	Result(Y) :- H(X,Y).
Cross-product:	Result(X,Y,U,V) :- R(X,Y), S(lJ,V).
Set-difference:	Result(X,Y) :- R(X,Y), NOT S(U,V).
lJnion:	H.esult(X,Y) :- R(X,Y).
	Result(X,Y) :- S(X,Y).

We conclude ()ur discussion of stratification l>y noting that SQL:1999 requires prograrns to be stratified. rrhe stratified Big/Sruall prograrn is shown below in SQL:1999 notation, with a final additional selection on Big2:

SQL:1999 and Datalog Queries: A Datalog rule is **linear** recursive if the body contains at Illost one occurrence of any table that depends on the table in the head of the rule. A linear recursive program contains only linear recursive rules. All linear recursive Datalog progranls can be expressed using the recursive features of SQL:1999. IIowever, these features are not in Core SQL.

```
WITH
Big2(Part) AS
        (SELECT   A1.Part FROM Assernbly Al WHERE Qty > 2)
Srnall2(Part) AS
        ((SELECT- A2.Part FROM Assernbly A2)
        EXCEPT
        (SELECT   Bl.Part fron1 Big2 Bl))

SELECT * FROM Big2 B2
```

24.4 FROM DATALOG TO SQL

To support recursive queries in SQL,we lllust take into account the features of SQL that are not found in Datalog. Two central SQL features rnissing in Datalog are (1) SQL treats tables as *multisets* of tuples, rather than sets, and (2) SQL pennits grouping and aggregate operations.

The rnultiset selnantics of SQL queries can be preserved if we do not check for duplicates after applying rules. Every relation instance, including instances of the recursively defined tables, is a lllultiset. rrhe nurnber of occurrences of a tuple in a relation is equal to the nurnber of distinct inferences that generate this tuple.

The second point can be addressed by extending Data.logwith grouping and aggregation operations. Tlhis rnust be done\vith rnultiset sernantics in rnind, as we now illustrate. Consider the following prograrn:

```
NumPartsCPart, SUM((Qty))) :- AssemblyCPart, Subpart, Qty).
```

This prograrn is equivalent to the SQL query

```
SELECT    A.Part, SUM (A.Qty)
FROM      Assernbly A
GROUP BY  A.Part
```

The angular brackets (...) notation was introduced in the LDL deductive system, one of the pioneering deductive database prototypes developed at MCC in the late 19808. We use it to dell0te *multiset generation*, or the creation of rnultiset-values. In principle, the rule definil1gNurnParts is evaluated by first creating the telnporary relation SHOWN in Figure 24.6. We create the ternporary relation by sorting on the *part* attribute (which appears on the left side of the rule, along with the (...) terrn) and collecting the I11ultiset of *qty* values for each *part* value. We then apply the SUM aggregate to each lllultiset-value in the second colu111n to obtain the answer, which is shown in Figure 24.7.

part	⟨*qty*⟩
trike	{3,1}
frarne	{1,1}
wheel	{2,1}
tire	{1,1}

Figure 24.6 Temporary Relation

part	[SUM(⟨*qty*⟩
trike	4
fran1e	2
wheel	3
tire	2

Figure 24.7 The Tuples in NumParts

The telnporary relation shown in Figure 24.6 need not be materialized to cornpute NurnParts; for exalllplc, SUM can be applied on-the-fly or Assenlbly can sirnply be sorted and aggregated as described in Section 14.6.

The use of grouping and aggregation, like negation, causes cOlnplicatiol1s when applied to a partially cOlnputed relation. rrhe difficulty is overcorne by adopting the sarne solution used for negation, stratification. Consider the following prograrn:[1]

```
TotParts(Part, Subpart, SUM«(Qty))) :- BOM(Part, Subpart, Qty).
BOM(Part, Subpart, Qty) :- Assembly(Part, Subpart, Qty).
BOM(Part, Subpart, Qty) :- Assembly(Part, Part2, Qty2),
        BOM(Part2, Subpart, Qty3), Qty=Qty2*Qty3.
```

The idea is to count the l111rnber of copies of Subpart for each Part. By aggregating over BOM rather than Assembly, we count subparts at any level in the hierarchy instead of just irnrnediate subparts. This prograrn is a version of a vvell-known problcrl1 called *Bill-of-Materials* and variants of it are probably the lnost widely used recursive queries in practice.

'rhe irnportant point to note in this exarnple is that we Inust wait until the relation BC)]VI has been cornpletely evaluated before we apply the rrotParts rule. Otherwise, \ve obta.in incornplete counts. This situation is analogous to theprobler11 we faced 'with negation; we have to evaluate the negated relation

[1] The reader should write this in SQL:1999 syntax, as a simple exercise.

SQL:1999 Cycle **Detection:** Safe Datalog queries that do not use arithmetic operations have finite answers and the fixpoint evaluation is guaranteed to halt. Unfortunately, recursive SQL queries may have infinite answer sets and query evaluation may not halt. There are two independent reasons for this: (1) the use of arithmetic operations to generate data values that are not stored in input tables of a query, and (2) multiset semantics for rule applications; intuitively, problems arise from *cycles* in the data. (To see this, consider the Components program on the Assembly instance shown in Figure 24.1 plus the tuple ⟨*tube, wheel,* 1⟩.) SQL:1999 provides special constructs to check for such cycles.

completely before applying a rule that involves the use of NOT. If a program is stratified with respect to uses of ⟨...⟩ as well as NOT, stratified fixpoint evaluation gives us meaningful results.

There are two further aspects to this example. First, we must understand the cardinality of each tuple in BOM, based on the multiset semantics for rule application. Second, we must understand the cardinality of the multiset of *Qty* values for each ⟨*Part, Subpart*⟩ group in TotParts.

part	*subpart*	*qty*
trike	frame	1
trike	seat	1
frame	seat	1
frame	pedal	2
seat	cover	1

Figure 24.8 Another Instance of Assembly

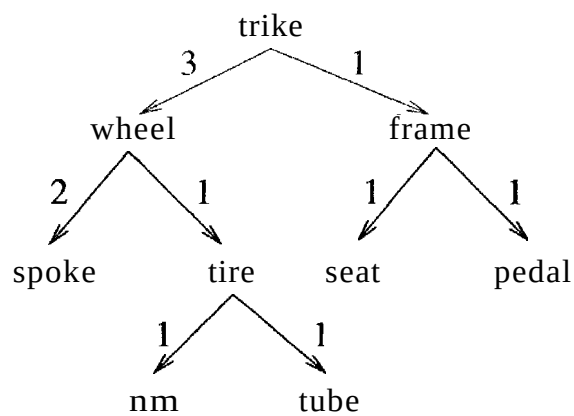


Figure 24.9 Assembly Instance Seen as a Graph

We illustrate these two points using the instance of Assembly shown in Figures 24.8 and 24.9. Applying the first BOM rule, we add (one copy of) every tuple in Assembly to BOM. Applying the second BOM rule, we add the following four tuples to BOM: ⟨*trike, seat,* 1⟩, ⟨*trike, pedal,* 2⟩, ⟨*trike, cover,* 1⟩, and ⟨*frame, cover,* 1⟩. Observe that the tuple ⟨*trike, seat,* 1⟩ was already in *BOM* because it was generated by applying the first rule; therefore, multiset semantics for rule application gives us two copies of this tuple. Applying the second BOM rule on the new tuples, we generate the tuple ⟨*trike, cover, 1*⟩ (using the tuple ⟨*frame, cover,* 1⟩ for Basic in the body of the rule): this is our second copy of the tuple. Applying the second rule again on this tuple does not generate any

tuples, and the cOInputation of the BOM relation is now cOlnplete. The BaM instance at this stage is sho\vn in Figure 24.10.

part	*subpart*	*qty*
trike	frarne	1
trike	seat	1
frame	seat	1
fraIne	pedal	2
seat	cover	1
trike	seat	1
trike	pedal	2
trike	cover	1
frame	cover	1
trike	cover	1

Figure 24.10 Instance of *BON!* Table

part	*subpart*	*qty*
trike	fraIlle	{1}
trike	seat	{1,1}
trike	cover	{1,1}
trike	pedal	{2}
fram.e	seat	{1}
frame	pedal	{2}
seat	cover	{I}
frame	cover	{I}

Figure 24.11 Temporary Relation

Multiset grouping on this instance yields the temporary relation instance shown in Figure 24.11. (This step is only conceptual; the aggregation can be done on the fly without materializing this terllporary relation.) Applying SUM to the rllultisets in the third column of this temporary relation gives us the instance for TotParts.

24.5 EVALUATING RECURSIVE QUERIES

rrhe evaluation of recursive queries has been widely studied. While all the problems of evaluating nonrecursive queries continue to be present, the newly introduced fixpoint operation creates additional difficulties. A straightforward approach to evaluating recursive queries is to cornpute the fixpoint by repeatedly applying the rules as illustrated in Section 24.1.1. One application of all the prograrn rules is caIled an iteration; we perfonn as rnany iterations as necessary to reach the least fixpoint. This approach has two rnain disadvantages:

- Repeated Inferences: As Figures 24.3 and 24.4 illustrate, inferences are repeated across iterations. That is, the sarne tuple is inferred repeatedly in *the same way*, using the same rule and the same tuples for tables in the body of the rule.
- Unnecessary Inferences: Suppose we want to find the components of only a *wheel*. Cornputing the entire Cornponents table is wasteful and does not take advantage of inforrnation in the query.

In this section, we discuss how each of these difficulties can be overcorne. We consider only Datalog prograins without negation.

24.5.1 Fixpoint Evaluation without Repeated Inferences

COlnputing the fixpoint by repeatedly applying all rules is called Naive fixpoint evaluation. Naive evaluation is guaranteed to cornpute the least fixpoint, but every application of a rule repeats all inferences lllade by earlier applications of this rule. We illustrate this point using the following rule:

```
Components(Part, Subpart) :- Assembly(Part, Part2, Qty),
                             Components(Part2, Subpart).
```

When this rule is applied for the first time, after applying the first rule defining Components, the Components table contains the projection of Assembly on the first two fields. Using these Components tuples in the body of the rule, we generate the tuples shown in Figure 24.3. For example, the tuple *(wheel, rim)* is generated through the following inference:

```
Components(wheel, rim) :- Assembly(wheel, tire, 1),
                          Components(tire, rim).
```

When this rule is applied a second time, the Components table contains the tuples shown in Figure 24.3 in addition to the tuples that it contained before the first application. Using the Components tuples shown in Figure 24.3 leads to new inferences; for example,

```
Components(trike, rim) :- Assembly(trike, wheel, 3),
                          Components(wheel, rim).
```

However, every inference carried out in the first application of this rule is also repeated in the second application of the rule, since all the Assernbly and Cornponents tuples used in the first rule application are considered again. For exarnple, the inference of *(wheel, rim⟩* shown above is repeated in the second application of this rule.

The solution to this repetition of inferences consists of rernelnbering which inferences were carried out in earlier rule applications and not carrying theln out again. We can 'remember' previously executed inferences efficiently by sirnply keeping track of which COlnponents tuples were generated for the first time in the rnost recent application of the recursive rule. Suppose we keep track by introducing a new relation called *delta_Components* and storing just the newly generated Cornponents tuples in it. Now, we can use only the tuples

in *delta_Components* in the next application of the recursive rule; any inference using other COIuponents tuples should have been carried out in earlier rule applications.

This refincrllcnt of fixpoint evaluation is called Seminaive fixpoint evaluation.Let us trace Serninaive fixpoint evaluation on our exarllple program. The first application of the recursive rule produces the Cornponents tuples shown in Figure 24.3, just like Naive fixpoint evaluation, and these tuples are placed in *delta_Components*. In the second application, however, only *delta_Components* tuples are considered, which rneans that only the following inferences are carried out in the second application of the recursive rule:

```
Components(trike, rim)  :- Assembly(trike, wheel, 3),
                           delta_Components(wheel, rim).
Components(trike, tube) :-Assembly(trike, wheel, 3),
                           delta_Components(wheel, tube).
```

Next, the bookkeeping relation *delta_Cornponents* is updated to contain just these two Cornponents tuples. In the third application of the recursive rule, only these two *delta_Cornponents* tuples are considered and therefore no additional inferences can be nlade. The fixpoint of Cornponents has been reached.

To irnplernent Serninaive fixpoint evaluation for general Datalog prograrns, we apply all the recursive rules in a prograrll together in an iteration. Iterative application of all recursive rules is repeated until no new tuples are generated in SOHle iteration. To surnrnarize how Serninaive fixpoint evaluation is carried out, there are two irnportant differences with respect to Naive fixpoint evaluation:

- We rnaintain a *delta* version of every recursive predicate to keep track of the tuples generated for this predicate in the Inost recent iteration; for example, *delta_Cornponents* for COHlponents. The *delta* versions are updated at the end of each iteration.
- The original prograrn rules are re\vritten to ensure that every inference uses at least one *delta* tuple; that is, one tuple that\vas not kno\vn before the previous iteration. This property guarantees that the inference could not have been carried out in earlier iterations.

We do lI0t discuss details of Serninaive fixpoint evaluation (such as the a.lgoritlun for rewriting prograrnl rules to ensure the use of a *delta* tuple in each inference).

24.5.2 Pushing Selections to Avoid Irrelevant Inferences

Consider a nonrecursive view definition. If we want only those tuples in the view that satisfy an additional selection condition, the selection can be added to the plan as a final operation, and the relational algebra transforlnations for conunuting selections with other relational operators allow us to 'push' the selection ahead of rnore expensive operations such as cross-products (;l,nd joins. In effect, we restrict the cornputation by utilizing selections in the query specification. The problerIl is rnore cOlnplicated for recursively defined queries.

We use the following progranl as an exarnple in this section:

```
8ameLevel(81 , 82)     Assembly(P1, 81, Q1),
                       Assembly(Pl, 82, Q2),
8ameLevel(81 , 82)     Assembly(Pl, 81, Qi),
                       8ameLevel(Pl, P2), Assembly(P2, 82, Q2).
```

Consider the tree representation of Assernbly tuples illustrated in Figure 24.2. 1"here is a tuple $\langle S1, S2\rangle$ in SarneLevel if there is a path froln 81 to 82 that goes up a certain nUlnber of edges in the tree and then CaInes down the saIne nurnber of edges.

Suppose we want to find all SalneLevel tuples with the first field equal to *spoke.* Since SalneLevel tuples can be used to COlupute other SarneLevel tuples, we cannot just cornpute those tuples with *spoke* in the first field. For exa.rnple, the tuple $\langle wheel, frarne\rangle$ in SarneLevel allows us to infer a SarneLevel tuple with *spoke* in the first field:

```
8ameLevel(spoke, seat) '-  Assembly(wheel, spoke, 2),
                           8ameLevel(wheel , frame),
                           Assembly(frame, seat, i),
```

Intuitively, we have to conlpute all SarneLevel tllpleswhose first field conta,ins a value on the path froln *spoke* to the root in Figure 24.2. Each such tuple has the potential to contribute to answers for the given query. On the other hand, cornputing the entire SameLevel table is wasteful; for exarnple, the SalneLevel tuple $\langle tire, seat\rangle$ cannot be used to infer any answer to the given query (or, indeed, to infer any tuple that can in turn be used to infer an answer tuple). We define a new table, \which we call l\1agic_SaIneLevel, such that each t11ple in this table identifies a value m for which we have to cornpute all SarneLevel tuples with m in the first colulun to answer the given query:

```
Magic_SameLevel(Pi) :- Magic_SameLevel(81), Assembly(P1, 81, Ql).
Magic_SameLevel(spoke) '-
```

Consider the tuples in lVlagic_SanleLevel. Obviously we have *(spoke)*. Using this lVlagic_SalneLevel tuple and the Assernbly tuple $\langle wheel, spoke, 2\rangle$, we can infer that the tuple $\langle wheel\rangle$ is in Magic_SameLevel. lJsing this tuple aJld the Assernbly tuple $\langle trike, 'wheel, 3\rangle$, we can infer that the tuple $\langle trike\rangle$ is in Nlagic_SarneLevel. Thus, Magic_SameLevel contains each node that is on the path frorn *spoke* to the root in Figure 24.2. The Magic_SarneLcvel table can be llsed as a filter to restrict the computation:

```
5ameLevel(51, 52) :- Magic_5ameLevel(51),
        Assembly(P1, 51, Q1), Assembly(P2, 52, Q2).
5ameLevel(51, 52) :- Magic_SameLevel(S1), Assembly(Pl, 51, Ql),
        5ameLevel(Pl, P2), Assembly(P2, 52, Q2).
```

These rules together with the rules defining rvlagic_SarneLevel give us a progranl for cornputing all SanleLevel tuples with *spoke* in the first column. Notice that the new progranl depends on the query constant *spoke* only in the second rule defining lVlagic_SameLevel. Therefore, the program for cornputing all SameLevel tuples with *seat* in the first column, for instance, is identical except that the second Magic_SarneLevel rule is

```
Magic_5ameLevel(seat) :- .
```

The nurnber of inferences rnade llsing the Magic program can be far fewer than the nurnber of inferences nlade using the original progranl, depending on just how rnuch the selection in the query restricts the cornputation.

24.5.3 The Magic Sets Algorithm

We illustrated the intuition behind the Magic Sets algorithrn on the SarneLevel prograrn, which contains just one output relation and one recursive rule.

The intuition behind the rewriting is that the rows in the Magic tables correspond to the subqueries whose answers are relevant to the original query. By evaluating the rewritten prograrn instead of the original prograrn, we can restrict cornputation by intuitively pushing the selection condition in the query into the recursion.

rIhe algorithrn, however, can be applied to any Datalog prograrn. The input to the algorithrn consists of the prograrn and a **query pattern**, which is a relation we want to query plus the fields for which a query will provide constants. The output of the algorithrn is a rewritten prograrn.

The Magic Sets program rewriting algorithrn can be surnrnarized as follows:

1. Generate the Adorned Prograln: In this step, the programl is re\vritten to l1lake the pattern of queries and subqueries explicit.
2. Add Magic Filters: IVlodify each rule in the Adorned Prograrn by adding a IVlagic condition to the body that acts as a filter on the set of tuples generated by this rule.
3. Define the Magic Tables: We create new rules to define the Magic tables. Intuitively, froIll each occurrence of a table R in the body of an Adorned PrograIu rule, we obtain a rule defining the table Magic_R.

When a query is posed, we add the corresponding Magic tuple to the rewritten prograrl1 and evaluate the least fixpoint of the prograrIl (using Serninaive evaluation).

We rernark that the Magic Sets algorithrll has turned out to be quite effective for cornputing correlated nested SQL queries, even if there is no recursion, and is used for this purpose in rnany cornrnercial DBlVISs, even systenls that do not currently support recursive queries.

We now describe the three steps in the Magic Sets algorithrIl using the SarneLevel program as a running exalllple.

Adorned Program

We consider the query pattern $SameLevel^{bf}$. Thus, given a value c, we want to cornpute all rows in $SarneLevel$ in which c appears in the first eolurnn. We generate the Adorned Prograrn P^{ad} frorn the given prograrn P by repeatedly generating adorned versions of rules in [J for every reachable query pattern, with the given query pattern as the only reachable pattern to begin with; additional reachable patterns are identified during the course of generating the A,dorned Prograrll as described next.

Consider a rule in? whose head contains the sarne table as sorne reachable pattern. rrhe adorned version of the rule depends on the order in \vhichwe consider the predicates in the body of the rule. To sirnplify our discussion, we assurne that this is always left-to-right. First, we replace the head of the rule with the rnatching query pattern. After this step, the recursive SarneLevel rule looks like this:

$SameLevel^{bf}$(S1, 82) :- Assembly(P1, 81, Q1),
8ameLevel(P1, P2), Assembly(P2, 82, Q2).

Next, we proceed left-to-right in the l)ody of the rule until we encounter the first recursive predicate. All cohullns that contain a constant or a variable that

appears to the left are marked *b* (for *bound*) and the rest are rnar,ked *f* (for *free*) in the query pattern for this occurrence of the predicate. We add this pattern to the set of reachal)le patterns and Inodify the rule accordingly:

$SameLevel^{bf}$(S1, S2) :- Assembly(Pl, 81, Ql),
$SameLevel^{bf}$(Pi, P2), AssemblyCP2, 82, Q2).

If there are additional occurrences of recursive predicates in the body of the recursive rule, we continue (adding the query patterns to the reachable set and rllodifying the rule). (()f course, in linear recursive prograins, there is at illOSt one occurrence of a recursive predicate in a rule body.)

\Ve repeat this until we have generated the adorned version of every rule in *P* for every reachable query pattern that contains the same table as the head of the rule. The result is the Adorned Program *pad*, which, in our example, is

$SameLeveZ^{bf}$C81, 82) :- AssemblyCP1, 81, Ql),
AssemblyCP1, 82, Q2).
$SameLevel^{bf}$(81, 82) :- AssemblyCP1, 81, Q1),
$SameLeveZ^{bf}$CP1, P2), AssemblyCP2, 82, Q2).

In our exarnple, there is only one reachable query pattern. In general, there can be several.[2]

Adding Magic **Filters**

Every rule in the Adorned Prograrn is rllodified by adding a 'nlagic filter' predicate to obtain the rewritten prograrn:

$Sarne\text{-}Level^{bf}$(81, 82) :- $Magic\_SameLevel^{bf}$(81),
Assembly(Pl, 81, Q1), Assembly(P2, 82, Q2).
$SarneLeveZ^{bf}$(S1, S2) :- $Magic_SameLevel^{bf}$CS1),
Assembly(Pl, 81, Ql), $8arneLevel^{bf}$(P1, P2),
Assembly(P2, 82, Q2).

The filter predicate is a copy of the head of the rule, 'with 'IVlagic' as a prefix for the table nàme and the variables in colllrnns corresponding to *free* deleted, as illustrated in these two rules.

[2] As an example, consider a variant of the SameLevel program in which the variables PI and *P*2 are interchanged in the body of the recursive rule (Exercise 24.5)

Defining Magic Filter Tables

Consider the Adorned Prograrl1 after every rule has been rnodified as described. FrorH each occurrence 0 of a recursive predicate in the body of a rule in this rIlodified prograrl1, we generate a rule that defines a Magic predicate. T'he algorithrl1 for generating this rule is as fo11o\:vs: (1) Delete everything to the right of occurrence () in the body of the rllodified rule. (2) Add the prefix 'Magic' and delete the free colulnns of (). (3) Move 0, with these changes, into the head of the rule.

From the recursive rule in our example, after steps (1) and (2) we get:

$$Sam,eLevel^{bf}(\texttt{S1},\ 82)\ \texttt{:-}\ Magic\_SarneLevel^{bf}(\texttt{S1}),\ \text{Assembly(P1, S1, Q1)},\ Magic\_SameLevel^{bf}(\texttt{P1})\ .$$

After step (3), we get:

$$Magic\_SameLevel^{bf}(\texttt{P1})\ \texttt{:-}\ Magic\_SameLevel^{bf}(\texttt{S1}),\ \text{Assembly(Pl, S1, Q1)}.$$

The query itself generates a row in the corresponding Magic table, for exarnple, $Magic\_SarneLevel^{bf}$(seat).

24.6 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- Describe *Datalog* prograrllS. lJse an exarnple Datalog prograrn to explain why it is not possible to write recursive rules in SQL-92. (Section 24.1)
- Define the terrns *rnodel* and *least model.* What can you say about least rnodels for Datalog prograrns? Why is this approach to defining the meaning of a Datalog prograrll called *declarative?* (Section 24.2.1)
- Define the tenns *.fi:rpoint* and *least ji:Epoint.* \iVhat can you say about least fixpoints for IJatalog prograrlls? vVhy is this approach to defining the rneaning of a Datalog prograrIl said to be *operational?* (Section 24.2.2)
- \What is a *safe* prograrn? Why is this property irIlportant? What is *range-restriction* and how does it ensure safety'? (Section 24.2.3)
- \What is the connection between least lnodels and least fixpoints for Datalog prograrIls? (Section 24.2.4)

- Explain why prograIIls with negation rnay not have a least model or least fixpoint. Extend the definition of *Tange-Testriction* to prograrns with negation. (Section 24.3)
- What is a *stratified* prograIn? How does stratification address the probleln of identifying a desired fixpoint? Show how every relational algebra query can be \vritten as a stratified Datalog prograrIl. (Section 24.3.1)
- Two important aspects of SQL, *multiset table8* and *aggr'egation 'with grouping,* are rnissing in Datalog. How can we extend Datalog to support these features? Discuss the interaction of these two new features and the need for stratification of aggregation. (Section 24.4)
- Define the terms *infeTence* and *iteration.* What are the two main challenges in efficient evaluation of recursive Datalog programs? (Section 24.5)
- Describe *Sem,inaive fixpoint evaluation* and explain how it avoids repeated inferences. (Section 24.5.1)
- Describe the *Magic Sets* program transformation and explain how it avoids unnecessary inferences. (Sections 24.5.2 and 24.5.3)

EXERCISES

Exercise 24.1 Consider the Flights relation:

Flights(fino: integer, *from:* string, *to:* string, *distance:* integer, *departs:* time, *arrives:* time)

Write the following queries in Datalog and SQL:1999 syntax:

1. Find the *fino* of all flights that depart from Madison.
2. Find the *flno* of all flights that leave Chicago after Flight 101 arrives in Chicago and no later than 1 hour after.
3. Find the *fino* of all flights that do not depart from Madison.
4. Find aJI cities reachable frOlll Madison through a series of one or 1I10re connecting flights.
5. Find all cities reachable from IVladison through a chain of one or rnore connecting flights, with no 1I1Ore than 1 hour spent on any connection. (That is, every connecting flight must depart within an hour of the arrival of the previous flight in the chain.)
6. Find the shortest tilne to fly frOl11 Madison to Madras, using a chain of one or 1nore connecting flights.
7. Find the *Jlno* of all flights that do not depart [1'0111 Madison or a city that is reacha.ble from Madison through a chain of flights.

Exercise 24.2 Consider the definition of Cornponents in Section 24.1.1. Suppose that the second rule is replaced by

```
Components(Part, Subpart) :-Components(Part, Part2),
                            Components(Part2, Subpart).
```

1. If the rnodified program is evaluated on the ASS€lnbly relation in Figure 24.1, how Inany iterations does Naive fixpoint evaluation take and what COlllponents facts are generated in each iteration?
2. Extend the given instance of Asselllbly so that Naive fixpoint iteration takes two rnore iterations.
3. Write this program in SQL:1999 syntax, using the WITH clause.
4. Write a prograni in Datalog syntax to find the part with the Inost distinct subparts; if several parts have the saIne Inaxinlllm number of subparts, your query should return all these parts.
5. How would your answer to the previous part be changed if you also wanted to list the number of subparts for the part with the Inost distinct subparts?
6. Rewrite your answers to the previous two parts in SQL:1999 syntax.
7. Suppose that you want to find the part with the rnost subparts, taking into account the quantity of each subpart used in a part, how would you rllodify the COlnponents program? *(Hint:* To write such a query you reason about the nuruber of inferences of a fact. For this, you have to rely on SQL's nlaintaining as many copies of each fact as the nurnber of inferences of that fact and take into account the properties of Seulinaive evaluation.)

Exercise 24.3 Consider the definition of Components in Exercise 24.2. Suppose that the recursive rule is rewritten as follows for Seminaive fixpoint evaluation:

```
Components(Part, Subpart) :- deLta_Components(Part, Part2, Qty),
                             deLta_Components(Part2, Subpart).
```

1. At the end of an iteration, what steps Illust be taken to update *delta_Cornponents* to contain just the new tuples generated in this iteration? Can you suggest an index on Cornponents that Inight help to make this faster?
2. Even if the *delta* relation is correctly updated, fixpoint evaluation using the preceding rule does not always produce all answers. Show an instance of Assembly that illustrates the probleru.
3. Can you suggest a way to rewrite the recursive rule in tenns of *delta_*Components so that Scrninaive fixpoint evaluation always produces all answers and no inferences are repeated across iterations?
4. Show how your version of the rewritten prograrn perfonns on the example instaJICe of Assernbly that you used to illustrate the problern with the gi"ven rewriting of the recursive rule.

Exercise 24.4 Consider the definition of SarneLevel In Section 24.5.2 and the Assernbly instance shown in Figure 24.1.

1. Rewrite the recursive rule for Seminaive fixpoint evaluation and show ho\v Serninaive evaluation proceeds.
2. Consider the rules defining the relation Magic, with *spoke* as the query constant. For Sernil1aive evaluation of the 'Magic' version of the SarneLevel prognllu, all tuples in Magic are cornputed first. Show how 8erninaive evaluation of the Magic relation proceeds.

3. After the Magic relation is cmnputed, it can be treated as a fixed database relation, just like Assembly, in the Senlinaive fixpoint evaluation of the rules defining SameLevel in the 'Magic' version of the prograrn. Rewrite the recursive rule for Selninaive evaluation and show how Scrninaive evaluation of these rules proceeds.

Exercise 24.5 Consider the definition of SanleLevel in Section 24.5.2 and a query in which the first argulnent is bound. Suppose that the recursive rule is rc\vritten as follows, leading to rnultiple binding patterns in the adorned program:

```
8ameLevel(81, S2) :- Assembly(Pl, 81, Ql),
                     Assembly(P1, 82, Q2).
8ameLevel(81, S2) :- Assembly(Pl, S1, Ql),
                     SameLevel(P2, P1), Assembly(P2, S2, Q2).
```

1. Show the adorned progranl.
2. Show the Magic program.
3. Show the Magic program after applying Seminaive rewriting.
4. Construct an example instance of Assenlbly such that the evaluating the optirnized prograrn generates less than 1% of the facts generated by evaluating the original prograrn (and finally selecting the query result).

Exercise 24.6 Again, consider the definition of SameLevel in Section 24.5.2 and a query in which the first argurnent is bound. Suppose that the recursive rule is rewritten as follows:

```
SameLevel(Sl, 82) :- Assembly(Pl, 81, Ql),
                     Assembly(Pl, S2, Q2).
SameLevel(Sl, 82) :- Assembly(P1, S1, Ql),
                     SameLevel(P1, Rl), SameLevel(Rl, P2), Assembly(P2, S2, Q2).
```

1. Show the adorned program.
2. Show the I\!Iagic prograln.
3. Show the Magic prograul after applying Serninaive rewriting.
4. Construct an exarnple instance of Asselnbly such that the evaluating the optimized pro-granl generates less than 1% of the facts generated by evaluating the original prograrnl (and finally selecting the query result).

BIBLIOGRAPHIC NOTES

The use of logic as a query language is discussed in several papers [296, 537], "which arose out of influential workshops. Good textbook discussions of deductive databases can be found in [747, 3, 143, 794, 503]. [614] is a recent survey article that provides an overview and covers the rnajor prototypes in the area, including LI)L [177], Glue-Nail! [214, 549] EKS-Vl [758], Aditi [615], Coral [612], LOLA [804], and XSB [644].

The fixpoint sernantics of logic programs (and deductive databases as a special case) is presented in [751], which also shows equivalence of the fixpoint seInantics to a *least-model* semantics. The use of stratification to give a natural sernantics to prograrns with negation was developed independently in [37, 154, 559,752].

Efficient evaluation of deductive database queries has been widely studied, and [58J is a survey and cOlnparison of several early techniques; [611] is a more recent survey_ Serninaive fixpoint evaluation was independently proposed several tiInes; a good treatment appears in [54]. r1'he Magic Sets technique is proposed in [57] and generalized to cover all deductive database queries without negation in [77]. The Alexander rnethod [631] was independently developed and is equivalent to a variant of Magic Sets called *Supplementary Magic Sets* in [77]. [553] shows how Magic Sets offers significant perfonnance benefits even for nonrecursive SQL queries. [673] describes a version of Magic Sets designed for SQL queries with correlation, and its irnplernentation in the Starbufst systern (which led to its iInplenlentation in IBNI's DB2 DBNIS). [670] discusses how Magic Sets can be incorporated into a Systenl R style cost-based optimization framework. The Magic Sets technique is extended to prograIIls with stratified negation in [53, 76]_ [121] cOlnpares Magic Sets with top-do\vn evaluation strategies derived froIn Prolog.

[642] develops a prograrn rewriting technique related to Magic Sets called *Magic Counting*. Other related methods that are not based on progranl rewriting but rather on fun-tirne control strategies for evaluation include [226, 429, 756, 757]. The ideas in 1.226] have been developed further to design an *abstract rnachine* for logic prograll1 evaluation using tabling in [609, 727]; this is the basis for the XSB systell1 [644].

25

DATA WAREHOUSING AND DECISION SUPPORT

- Why are traditional DBIvISs inadequate for decision support?
- What is the multidimensional data nlOdel and what kinds of analysis does it facilitate?
- What SQL:1999 features support rnultidiInensional queries?
- How does SQL:1999 support analysis of sequences and trends?
- How are DBMSs being optimized to deliver early answers for interactive analysis?
- What kinds of index and file organizations do OLAP systerlls require?
- What is data warehousing and why is it irnportant for decision support?
- Why have rnaterialized views becorne iInportant?
- How can we efficiently Inaintain rnaterialized views?
- Key concepts: OLAP, rnultirnensional rnodel, dimensions, measures; roll-up, drill-clown, pivoting, cross-tabulation, CUBE; WINDOW queries, frames, order; top N queries, online aggregation; bitmap indexes, join indexes; data warehouses, extract, refresh, purge; rnaterialized views, incremental rnaintenancc, rnaintaining warehouse views

Notlling is lnore difficult, and therefore more precious, than to be able to decide.

. NCtI)oleon Bonaparte

Database luanagclncnt systerIls are widely used by organizations for rnaintaining data that docurnents their everyday operations. In applications that update such *operational data*, transactions typically rnake srna11 changes (for exalnple, adding a reservation or depositing a check) and a large nU111ber of transactions H1USt be reliably and efficiently processed. Such online transaction processing (OLTP) applications have driven the gruwth of the DBMS industry in the past three decades and will doubtless continue to be irnportant. DB1VISs have traditionally been optirnized extensively to perforn1 well in such applications.

H,ecently, ho\vever, organizations have increasingly crnphasized applications in which current and historical data is coruprehensively analyzed and explored, identifying useful trends and creating sununaries of the data, in order to support high-level decision rnaking. Such applications are referred to as decision support. Mainstream relational DBMS vendors have recognized the irnportance of this rnarket segment and are adding features to their products to support it. In particular, SQL has been extended with new constructs and novel indexing and query optirl1ization techniques are being added to support cornplex queries.

The use of views has gained rapidly in popularity because of their utility in applications involving cornplex data analysis. While queries on views can be answered by evaluating the view definition when the query is subrnitted, precornputing the view definition can rnake queries run Inuch faster. Carrying the r11otivation for preconlputed views one step further, organizations can consolidate inforrnation from several databases into a *data warehouse* by copying tables fror11 rnany sources into one location or rnaterializing a view defined over tables fr01n several sources. Data '\varehousing has becorne widespread, and Il1any specialized products are no\v available to create and rnanage warehouses of data frorH 1l1ultiple databases.

We begin this chapter with an overview of decision support in Section 25.1. We introduce the rnultirnensional rnodel of data in Section 25.2 and consider database design issues in 25.2.1. We discuss the rich class of queries that it naturally supports in Section 25.;3. We discuss how new SQL:1999 constructs allow us tel express rnultidilnensional queries in 25.3.1. In Section 25.4, we discuss SQL:1999 extensions that support queries over relations as ordered collections. We consider how to optimize for fast generation of initial answers in Sectioll 25.5. The rnany query language extensions required in the ()LA.P envirolllnentprornpted the developrnent of llC\V irnplcrncntation techniques; we discuss these in Section 25.6. In Section 25.7, \ve examine the issues involved in creating and rnaintaining a data \varehouse. FraIn a technical standpoint, a key issue is how to maintain warehouse inforrnation (replicated tables or views) ·when the llnderl.ying source infonnation changes. After covering the important role played byvic\vs in OLAP and warehousing irl Section 25.8, we consider maintenance of rnaterialized views in Sections 25.9 and 25.10.

25.1 INTRODUCTION TO DECISION SUPPORT

()rganizational decisioll rnaking requires a cOlnprehensive view of all aspects of an enterprise, so many organizations created consolidated data warehouses that contain data drawn frcHn several databases ll1atntained by different business units together with historical and summary inforInation.

The trend toward data warehousing is c0I11plelnented by an increased ernphasis on powerful analysis tools. Many characteristics of decision support queries make traditional SQL systenls inadequate:

- 1'he WHERE clause often contains rnany AND and OR conditions. As we saw in Section 14.2.3, OR conditions, in particular, are poorly handled in rnany relational DBMSs.
- Applications require extensive use of statistical functions, such as standard deviation, that are not supported in SQL-92. Therefore, SQL queries rnust frequently be ernbedded in a host language program.
- Many queries involve conditions over time or require aggregating over time periods. SQL-92 provides poor support for such time-series analysis.
- Users often need to pose several related queries. Since there is no convenient way to express these cOlnnlonly occurring families of queries, users have to write thern as a collection of independent queries, \vhich can be tedious. Further, the DBMS has no way to recognize and exploit optimization opportunities arising froln executing nlany related queries together.

Three broad classes of analysis tools are available. First, SOlne systerIls support a class of stylized queries that typically involve group-by and aggregation operators and provide excellent support for cOlnplex boolean conditions, statistical functions, and features for tilne-series analysis. Applications dominated by such queries are called online analytic processing (OLAP). 'These systerns support a querying style in which the data is best thought of as a rnultidilnensional array and are influenced by end-user tools, such as spreadsheets, in addition to database query languages.

Second, sorne DBMSs support traditional SQL-style queries but are designed to also support OLAP queries efficiently. Such systenls can be regarded as relational DBMSs optirnized for decision support applications. Many vendors of relational DBIVISs are currently enhancing their products in this direction and, over tilne, the distinction between specialized OLAP systerns and relational DBIVISs enhanced to support ()LAP queries is likely to dirninish.

The third class of analysis tools is rllotivated by the desire to find interesting or unexpected trends and patterns in large data sets rather than the conlplex

SQL:1999 and OLAP: In this chapter, we discuss a nUInber of features introduced in SQL:1999 to support OLAP. In order not to delay publication of the SQL:1999 standard, these features \vere actually added to the standard through an *amendment* called SQL/OLAP.

query characteristics just listed. In exploratory data analysis, although an analyst can recognize an :interesting pattern' when shown such a pattern, it is very difficult to fannulate a query that captures the essence of an interesting pattern. For exalnple, an analyst looking at credit-card usage histories Illay want to detect unusual activity indicating Inisuse of a lost or stolen card. A catalog lllerchant lnay want to look at custolner records to identify pro111ising custoiners for a new proillotion; this identification would depend on inccnIle level, buying patterns, delllonstrated interest areas, and so all. The alllount of data in Inany applications is too large to perrnit rnanual analysis or even traditional statistical analysis, and the goal of data mining is to support exploratory analysis over very large data sets. We discuss data rnining further in Chapter 26.

Clearly, evaluating OLAP or data rnining queries over globally distributed data is likely to be excruciatingly slow. Further, for such cOlnplex analysis, often statistical in nature, it is not essential that the IllOSt current version of the data be used. The natural solution is to create a centralized repository of all the data; that is, a data warehouse. Thus, the availability of a warehouse facilitates the application of ()LAP and data rnining tools and, conversely, the desire to apply such analysis tools is a strong 1Ilotivation for building a data warehouse.

25.2 OLAP: MULTIDIMENSIONAL DATA MODEL

aLAI' applications are dOlninated by ad hoc, cOlnplex queries. In SQL terllls, these are queries that involve group-by and aggregation operators. The natural way to think about typical ()LAP queries, ho\vever, is in tenns of a rnultidilnensinnal data rllodel. In this section, we present the rnultidirnensional data Illodel and corupare it with a relational representation of data. In subsequent sections, we describe ()LAP queries in. terrns of the rllultidirnensional data rnodel and consider some new irnplernentation techniques designed to support such queries.

In the rnultidirnensional data rnodel, the focus is on a collection of nurneric measures. Each Ineasure depends on a set of dirnensions. We use a running exarnple based on sales data. The measure attribute in our exarnple is *sales*. The dirnensions are Product, Location, and Tirne. Given a product, a location;

and a tinle, we have at I110st one associated sales value. If we identify a product by a unique identifier *pid* and, sirnilarly, identify location by *locid* and time by *timeid*, we can think of sales inforrnation as being arranged in a three-dirnensional array Sales. This array is shown in Figure 25.1; for clarity, we show only the values for a single *locid* value, *locid-* 1, which can be thought of as a slice orthogonal to the *lacid* axis.

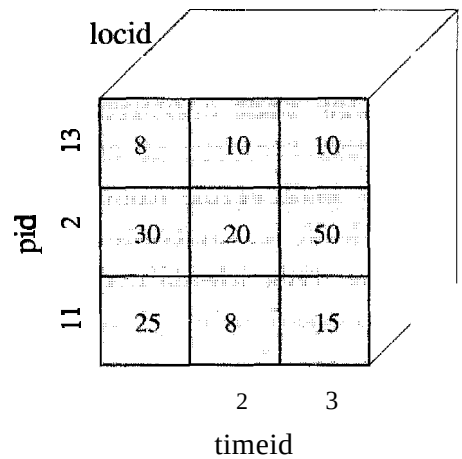


Figure 25.1 Sales: A Multidimensional Dataset

'This view of data as a multiclhnensional array is readily generalized to rnore than three dirnensions. In OLAP applications, the bulk of the data can be represented in such a rnultidiInensional array. Indeed, some OLAP systerns actually store data in a rnultidiInensional array (of course, irnplenlented without the usual prograrnrning language asslunption that the entire array fits in rnelnory). 0 LAP systerns that use arrays to store rnultidirnensiona.l datasets are called nlultidimensional OLAP (MOLAP) systcrl1s.

The data in a 11lultidirnensional array can also be represented as a relation, as illustrated in Figure 25.2, which shows the same data as in Figure 25.1, with additional rows corresponding to the 'slice' *locid*= 2. Tlhis relation, which relates the dirnensions to the rneasure of interest, is called the fact table.

Now let us tun1 to dirnensions.Each dirnension can have a set of associated attributes. For exarnple, the Location dilTlension is identified by the *loc'id* attribute, which we used to identify a location in the Sales table. We aSSUlnc that it also has attributes *country*, *state*, and *city*. We further assurne that the Product dirnension has attributes *pname*, *category*, and *price* in additi(Jn to the identifier *pid*. 'rhe *category* of a product indicates its general nature; for exarnple, a product *pant* could have category value *apparel*. We assurne that the Time dirnension has attributes *date*, *week*, *month*, *quarter*, *year*, and *holiday_flag* in addition to the identifier *timeid*.

locid	*city*	*state*	*country*
1	Madison	WI	USA
2	Fresno	CA	USA
5	Chennai	TN	India

Locations

pid	*pname*	*category*	*price*
11	Lee Jeans	Apparel	25
12	Zord	Toys	18
13	Biro Pen	Stationery	2

Products

pid	*timeid*	*locid*	*sales*
11	1	1	25
11	2	1	8
11	3	1	15
12	1	1	30
12	2	1	20
12	3	1	50
13	1	1	8
13	2	1	10
13	3	1	10
11	1	2	35
11	2	2	22
11	3	2	10
12	1	2	26
12	2	2	45
12	3	2	20
13	1	2	20
13	2	2	40
13	3	2	5

Sales

Figure 25.2 Locations, Products, and Sales Represented as Helations

For each dinlension, the set of associated values can be structured as a hierarchy. For exarnple, cities belong to states, and states belong to countries. Dates belong to weeks and rIlonths, both 'weeks and 1110nths are contained in quarters, and quarters are contained in years. (Note that a week could span two rnonths; therefore, weeks are not contained in rnonths.) SCHne of the attributes of a diruension describe the position of a dirnensioll valuevvith respect to this underlying hierarchy of dirnensioll values. The hierarchies for the Product, Location, and Time hierarchies in our exarnple are sho\vn at the attribute level in Figure 25.3.

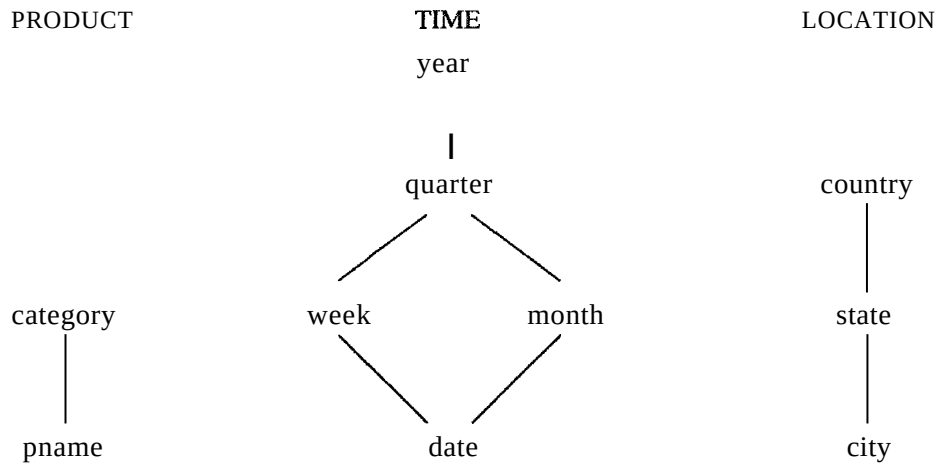


Figure 25.3 Dimension Hierarchies

Infonnation about dirnensions can also be represented as a collection of relations:

Locations(*locid:* integer, *city:* string, *state:* string, *country:* string)
Products(pid: integer, *pname:* string, *category:* string, *price:* real)
Times(*timeid:* integer, *date:* string, *week:* integer, *rnonth:* integer, *quarter:* integer, *year:* integer, *holiday_fiag:* boolean)

These relations arc luuch srnalIer than the fact table in a typical 0 I..lAP application; they are called the **dimension** tables. OLAP systcrl1s that store all inforrnation, including fact tables, as relations are called relational OLAP (ROLAP) systcrns.

The Tinlcs table illustrates the attention paid to the T'irne dirnension in typical OLAP applications. SQL's date and tirnestaulp data types are not adequate; to support slunrnarizations that reflect business operations, infonnation such as fiscal quarters, holiday status, and so on is rnaintained for each tirne value.

25.2.1 Multidimensional Database Design

Figure 25.4 shows the tables in our running sales example. It suggests a star, centered at the fact table Sales; such a cornbination of a fact table and dirncnsion tables is called a star schema. This schelna pattern is very COIIUIlon in databases designed for 0 LAP. The bulk of the data is typically in the fact table, which has no redundancy; it is usually in BCNF. In fact, to Ininimize the size of the fact table, dirnension identifiers (such as *p'id* and *timeid*) are systcrn-generated identifiers.

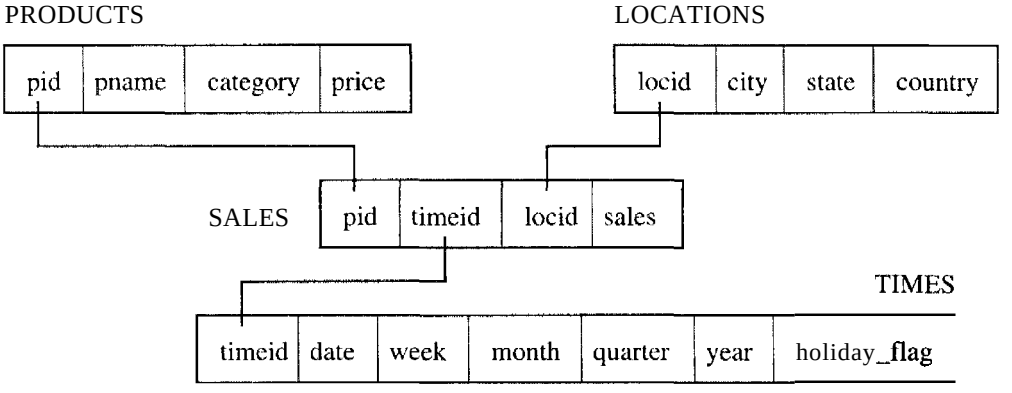


Figure 25.4 An Example of a Star Schema

Inforrnation about dinlension values is rnaintained in the dirnension tables. Di-111ension tables are usually not nonnalized. The rationale is that the dimension tables in a database used for OL,AP are static and update, insertion, and deletion anoillalies are not irnportant. Further, because the size of the database is dorninated by the fact table, the space saved by norrnalizing dilnension tables is negligible. Therefore, rnini111izing the cornputation tilllC for cOlllbining facts in the fact table with dirnension inforrnation is the rnain design criterion, which suggests that we avoid breaking a dirnension table into srnaller tables (which rnight lead to additional joins).

Snlall response tirnes for interactive querying are irnportant in OLAP, and rnost systerns support the Hlaterialization of SUrl1Inary tables (typically generated through queries using grouping). Ad hoc queries posed by users are answered using the original ta,bles along with precornputed surnrnaries. A very irnportant design issue is which sunnnary tables should be rnaterialized to achieve the best use of available rnerllory and answer cOHnI1only asked ad hoc queries with interactive response tirnes. In current OLAP systerns, deciding which surnnlary tables to rnaterialize rnay well be the Inost irnportant design decision.

Finally, new storage structures and indexing techniques have been developed to support ()LAP and they present the database designer \'lith additional physical

design choices. We cover BOIHe of these hnplclnentatiol1 techniques in Section 25.6.

25.3 MULTIDIMENSIONAL AGGREGATION QUERIES

Now that we have seen the rnulticliInensiol1alluoclel of data, let us consider how such data can be queried and rnanipulatecl. The operations supported by this Inodel are strongly influenced by end user tools such as spreadsheets. The goal is to give end users who are not SQL experts an intuitive and po\verful interface for cornnlon business-oriented analysis tasks. Users are expected to pose ad hoc queries directly, without relying on database application prograrrnners.

In this section, we asslllne that the user is working with a multidirnensional dataset and that each operation returns either a different presentation or a sunllnary; the underlying dataset is always available for the user to 1nanipulate, regardless of the level of detail at which it is currently viewed. In Section 25.3.1, we discuss how SQL:1999 provides constructs to express the kinds of queries presented in this section over tabular, relational data.

A very C01111non operation is aggregating a rneasure over one or 1nore dimensions. The following queries are typical:

- Find the total sales.
- Find total sales for each city.
- Find total sales for each state.

'These queries can be expressed as SQL queries over the fact and dirnension tables. When we aggregate a measure Oll one or rnore di1nensions, the aggregated measure depends on fewer diInensiolls than the original measure. For example, when we cornpute the total sales by city, the aggregated rneasure is *total sales* and it depends only on the Location di1nension,whereas the original *sales* rneasure depended on the Locatioll,Tirne, a,nd Product dirnensions.

Another use of aggregation is to SU1Ilrnarize at different levels of a dirnension hierarchy. If we are given total sales per city, we can aggregate 011 the Location dinlension to obtain sales per state. This operation is called roll-up in the OLAI' literature. The inverse of roll-up is drill-down: Given total sales by state, we can ask for a 1Ilore detailed presentation by drilling down on Location. We can ask for sales by city or just sales by city for a selected state (with sales presented on a per-state basis for the rernaining states, as before). We can also drill dowll on a diluension other than Location. For example, we can ask

for total sales for each product for each state, drilling down on the Prodnet diInension.

Another C0111l10n operation is pivoting. Consider a, tabular presentation of the Sales table. If we pivot it on the Location and Titne dirnensions, we obtain a table of total sales for each location for each tillle value. This infoI"luation can be presented as a two-dimensional chart in which the axes are labeled 'with location and time values; the entries in the chart correspond to the total sales for that location and time. Therefore, values that appear in colurnns of the original presentation becoIne labels of axes in the result presentation. The result of pivoting, called a cross-tabulation, is illustrated in Figure 25.5. Observe that in spreadsheet style, in addition to the total sales by year and state (taken together), we also have additional sunlillaries of sales by year and sales by state.

	WI	CA	Total
1995	63	81	144
1996	38	107	145
1997	75	35	110
Total	176	223	399

Figure 25.5 Cross-Tabulation of Sales by Year and State

Pivoting can also be used to change the dirnensions of the cross-tabulation; froIn a presentation of sales by year and state, we can obtain a presentation of sales by product and year.

Clearly, the OLAP frarnework rnakes it convenient to pose a broad class of queries. It also gives catchy naInes to sorne farniliar operations: Slicing a dataset arnonnts to an equality selection on one or rIlore dirnensions, possibly also with SC)lne dirnensions projected out. Dicing a dataset arl10unts to a range selection. These terrllS corne frcnl1 visuaJizing the effect of these operations on a cube or cross-tabulated representation of the data.

A Note on Statistical Databases

Many ()LAP concepts are present in earlier work on statistical databases (SDBs), which are database systerl1s designed to support statistical applications, although this connection has not been sufficiently recognized because of differences in application dornains and tern.linology. The rnultidirnensional data rllodel, 'with the notions of a rneasure associated with dirnensions and

classification hierarchies for dirncIlsion vahles, is also used in SDBs. OLAP operations such as roll-up and drill-down have counterparts in SDBs. Indeed, sorne irnplcrnentation techniques developed for OLAP are also applied to SDBs.

Nonetheless, some differences arise frorn the different dOlnains OLAP and SDBs were developed to support. For exarnple, SnBs are used in socioeconornic applications, where classification hierarchies and privacy issues are very ilnportant. This is reflected in the greater cornplexity of classification hierarchies in SDBs, along with issues such as potential breaches of privacy. (The privacy issue concerns whether a user with access to sUllunarized data can reconstruct the original, unsununarized data.) In contrast, OLAP has been ailned at business applications with large volulnes of data and efficient handling of very large datasets has received lnore attention than in the SDB literature.

25.3.1 ROLLUP and **CUBE** in SQL:1999

In this section, we discuss how lnany of the query capabilities of the rnultidi-111ensionalrlloclel are supported in SQL:1999. Typically, a single OLAP operation leads to several closely related SQL queries with aggregation and grouping. For exarnple, consider the cross-tabulation shown in Figure 25.5, which was obtained by pivoting the Sales table. To obtain the saIne inforrnation, we would issue the following queries:

```
SELECT    rr.year, 1.state, SUM (S.sales)
FROM      Sales S, T'irnes T, Locations L
WHERE     S.tirneid=T.tiIneid AND S.locid=L.locid
GROUP BY  T.year, 1.state
```

This query generates the entries in the body of the chart (outlined by the dark lines). The surllluary cohunn on the right is generated by the query:

```
SELECT    T.year, SUM (S.saJes)
FROM      Sales S1 Times T
WHERE     S.timeid= T.tunclC1
GROUP BY  T.year
```

The sunnnary ro\v at the bottorl1 is generated l)y the query:

```
SELECT    L.state, SUM (S.sales)
FROM      Sales S, Locations L
WHERE     S.locid=L.locicl
GROUP BY  L.state
```

The cumulative SUITI in the bottonl-right corner of the chart is produced by the query:

```
SELECT   SUM (S.sales)
FROM     Sales S, Locations L
WHERE    S.loc:id=L.locid
```

The exarnple cross-tabulation can be thought of as roll-up on the entire dataset (i.e., treating everything as one big group), on the Location dirnension, on the rrirne dirnensioIl, and on the Location and Tinle dinlensions together. Each roll-up corresponds to a single SQL query with grouping. In general, given a measure with k associated dirnensions, we can roll up on any subset of these k diInensions; so we have a total of 2^k such SQL queries.

Through high-level operations such as pivoting, users can generate lTlany of these 2^k SQL queries. R,ecognizing the cornrnonalities between these queries enables r110re efficient, coordinated COlTlputation of the set of queries.

SQL:1999 extends the GROUP BY construct to provide better support for roll-up and cross-tabulation queries. The GROUP BY clause with the CUBE keyword is equivalent to a collection of GROUP BY statenlents, with one GROUP BY statement for each subset of the k dirnensions.

Consider the following query:

```
SELECT   rr.year, L.state, SUM (S.sales)
FROM     Sales S, Tirnes T, Locations L
WHERE    S.tirneid=T'.tirneid AND S.1ocid=L.locid
GROUP BY CUBE (T.year, L.state)
```

The result of this query, shown in Figure 25.6, is just a tabular representation of the cross-tabulation in Figure 25.5.

SQL:1999 also provides variants of GROUP BY that enable cornputatioll of subsets of the cross-tabulation cornputed using GROUP BY CUBE. For example, we can replace the grouping clause in the previous query with

```
GROUP BY ROLLUP (T.year, L.state)
```

In contrast to GROUP BY CUBE, vile aggregate by an pairs of year and state values and by each year, and cornpute an overall SIIHI for the entire dataset (the last row in Figure 25.6), but we do not aggregate for each state value. The result is identical to that shown inF'igure 25.6, except that the rows with *null* in the *T.year* COhUIIII and non-*null* values in the *L.state* colurnn are not cornputed.

```
CUBE pid, locid, tirneid BY SUM Sales
```

T.year	L.state	SUM(S.sales)
1995	WI	63
1995	CA	81
1995	*null*	144
1996	WI	38
1996	CA	107
1996	*null*	145
1997	WI	75
1997	CA	35
1997	*null*	110
null	WI	176
null	CA	223
null	*null*	399

Figure 25.6 The Result of GROUP BY CUBE on Sales

rrhis query rolls up the table Sales on all eight subsets of the set {pid, locid, tirneid} (including the empty subset). It is equivalent to eight queries of the fonn

```
SELECT    SUM (S.sales)
FROM      Sales S
GROUP BY  grouping-list
```

The queries differ only in the *grouping-list,* which is sorne subset of the set {pid, locid, tirneid}. We can think of these eight queries as being arranged in a lattice, as shown in Figure 25.7. The result tuples at a node can be aggregated further to cornpute the result for any child of the node. This relationship between the queries arising in a CUBE can be exploited for efficient evaluation.

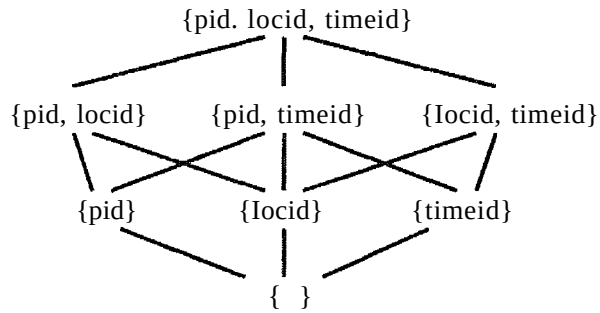


Figure 25.7 The Lattice of GROUP BY Queries ill a CUBE Query

25.4 WINDOW QUERIES IN SQL:1999

The time dirnension is very important in decision support and queries involving trend analysis have traditionally been difficult to express in SQL. To address this, SQL:1999 introduced a fundamental extension called a query window. Examples of queries that can be written using this extension, but are either difficult or iInpossible to write in SQL without it, include

1. Find total sales by rnonth.
2. Find total sales by rnonth for each city.
3. Find the percentage change in the total monthly sales for each product.
4. Find the top five products ranked by total sales.
5. Find the trailing *n* day moving average of sales. (For each day, we must compute the average daily sales over the preceding *n* days.)
6. Find the top five products ranked by cumulative sales, for every month over the past year.
7. Rank all products by total sales over the past year, and, for each product, print the difference in total sales relative to the product ranked behind it.

The first two queries can be expressed as SQL queries using GROUP BY over the fact and dinlension tables. The next two queries can be expressed too, but are quite complicated in SQL-92. The fifth query cannot be expressed in SQL-92 if *n* is to be a pararneter of the query. The last query cannot be expressed in SQL-92.

In this section, we discuss the features of SQL:1999 that allow us to express all these queries and, obviously, a rich class of sirnilar queries.

The rnain extension is the WINDOW clause, which intuitively identifies an ordered 'window' of rows 'around' each tuple in a table. This allows us to apply a rich collection of aggregate functions to the windovv of a row and extend the row with the results. For exarnple, we can associate the average sales over the past 3 days with every Sales tuple (each of which records 1 day's sales). This gives us a 3-day Illoving average of sales.

\While there is sorne sirnilarity to the GROUP BY and CUBE clauses, there are important differences as well. For exarnple, like the WINDOW operator, GROUP BY allows us to create partitions of rows and apply aggregate functions such as SUM to the rows in a partition. However, unlike WINDOW, there is a single output row per pa.rtition, rather than one output row for each row, and each partition is an unordered collection of rows.

We now illustrate the window concept through an exalnple:

```
SELECT L.state, T.month, AVG (S.sales) OVER W AS Inovavg
FROM   Sales S, Tinles T, Locations L
WHERE  S.tirneid=T.tirIleid AND S.1ocid=L.locid
WINDOW W AS (PARTITION BY L.state
        ORDER BY 'f.lnonth
        RANGE BETWEEN INTERVAL '1' MONTH PRECEDING
        AND INTERVAL '1' MONTH FOLLOWING)
```

The FROM and WHERE clauses are processed as usual to (conceptually) generate an interrnediate table, which we refer to as Ternp. Windows are created over the TeHIp relation.

There are three steps in defining a window. First, we define *part'it'ions* of the table, using the PARTITION BY clause. In the exarnple, partitions are based on the *L.8tate* colurnn. Partitions are similar to groups created with GROUP BY, but there is a very important difference in how they are processed. To understand the difference, observe that the SELECT clause contains a column, *T. month,* which is not used to define the partitions; different rows in a given partition could have different values in this colulun. Such a colurnn cannot appear in the SELECT clause in conjunction with grouping, but it is allowed for partitions. 'The reason is that there is one answer row for *each* row in a partition of Ternp, rather than just one answer row per partition. The window around a given row is used to COlnpute the aggregate functions in the corresponding answer row.

The second step in defining a \vindow is to specify the *ordering* of rows within a partition. We do this using the ORDER BY clause; in the exarnple, the rows within each partition are ordered by *T. month.*

The third step in window definition is to *frame* windo\vs; that is, to establish the boundaries of the window associated with each row in terrns of the ordering of rows within partitions. In the exalnple, the window for a row includes the row itself plus all rows whose rnonth value is within a Inonth before or after; therefore, a row whose *Tnonth* value is Jllne 2002 has a window containing all rows with *Tnonth* equal to May, June, or July 2002.

1'he answer row corresponding to a given row is constructed by first identifying its \vindo\v. Then, for each answer colurun defined using a window aggregate function, we cornpute the aggregate llsing the ro\vs in the window.

In our example, each row of Temp is essentially a ro\v of Sales, tagged with extra details (about the location and tirne dirnensions). There is one partition for ea.ch state and every ro\v of Ternp belongs to exactly one partition. Consider

a ro\v for a store in \Visconsin. The row states the sales for a given product, in that store, at a certain tirHe. The window for this row includes all rows that describe sales in \Visconsin within the previous or next Inonth and *movavg* is the average of sales (over all products) in \Visconsin \vithin this period.

We note that the ordering of rows within a partition for the purposes of window definition does not extend to the table of answer ro\vs. The ordering of answer rows is nondeterlninistic, unless, of course, we fetch therIl through (1, cursor and use ORDER BY to order the cursor's output.

25.4.1 Framing a Window

There are two distinct ways to fra1118 a window in SQL:1999. The exarnple query illustrated the RANGE construct, which defines a window based on the values in SOllle cohulln *(rnonth* in our exarnple). The ordering colu111n has to be a nU111eric type, a datetillle type, or an interval type since these are the only types for which addition and subtraction are defined.

The second approach is based on using the ordering directly and specifying how Illany rows before and after the given row are in its window. Thus, we could say

```
SELECT  L.state, T.rnonth, AVG (S.sales) OVER W AS Inovavg
FROM    Sales S, Times T, Locations L
WHERE   S.timeid=T.timeid AND S.locid=L.locid
WINDOW W AS (PARTITION BY L.state
             ORDER BY T.IIlonth
             ROWS BETWEEN 1 PRECEDING AND 1 FOLLOWING)
```

If there is exactly one row in Tenlp for each IIlonth, this is equivalent to the previous query. IIo\vever, if a given lnonth has no rows or lnultiple rows, the t\VO queries produce different results. In this case, the result of the second query is hard to understand because the \vindc)\vs for different rows do not align in a natural way.

The second approach is appropriate if, in terms of our example, there is exactly one 1'o\v per lllonth. C·eneralizing frOIrI this, it is also appropriate if there is exactly one row for every value in the sequence of ordering COhll1Ul values. ·UnJike the first approach, 'where the ordering has to be specified over a single (rullneric, datetime, or interval type) colurnn, the ordering can be based on a cornposite key.

We can also define \vindows that include all rows that are before a given row (UNBOUNDED PRECEDING) or all rO\VS after a given row (UNBOUNDED FOLLOWING) 'within the row's partition.

25.4.2 New Aggregate Functions

While the standard aggregate functions that apply to rnultisets of values (e.g., SUM, AVG) can be used in conjunction \vith Willdo\ving, there is a lleed for a new class of functions that operate on a *list* of values.

The RANK function returns the position of a row within its partition. If a partition has 15 rows, the first row (according to the ordering of rows in the window definition over this partition) has rank 1 and the last row has rank 15. The rank of intermediate rows depends on whether there are multiple (or no) rows for a given value of the order.ing colurnn.

Consider our running example. If the first row in the Wisconsin partition has the lllonth January 2002, and the second and third rows both have the rnonth February 2002, then their ranks are 1, 2, and 2, respectively. If the next row has rllonth March 2002 its rank is 4.

In contrast, the DENSE_.RANK function generates ranks without gaps. In our exalnple, the four rows are given ranks 1, 2, 2, and 3. The only change is in the fourth row, whose rank is now 3 rather than 4.

The PERCENT_RANK function gives a lneasure of the relative position of a row within a partition. It is defined as (RANK-1) divided by the Innnber of rows in the partition. CUME_DIST is sirnilar but based on actual position within the ordered partition rather than rank.

25.5 FINDING ANSWERS QUICKLY

A recent trend, fueled in part by the popularity of the Internet, is an ernphasis 011 queries for which a user wants only the first few, or the 'best' few, answers quickly. When users pose queries to a search engine such as AltaVista, they rarely look beyond the first or second page of results. If they do not find what they are looking for, they refine their query and resubrnit it. The same phen()lneuon occurs in decision support applications and scnne DBl\;1S products (e.g., DB2) already support extended SQL con.structs to specify sueh queries. A related trend is that, for cornplex queries, users would like to see an approximate answer quickly and then have it be continually refined, rather than \vait until the exact answer is available. We now discuss these two trends l)riefly.

25.5.1 Top N Queries

An analyst often wants to identify the top-selling handful of products, for ex-alnple. We can sort by sales for each product and return answers in this order. If we have a Inillion products and the analyst is interested only in the top 10, this straightforward evaluation strategy is clearly \vasteful. It is desirable for users to be able to explicitly indicate how rnany answers they want, rnaking it possible for the DBMS to optirnize execution. l-'he follo\ving exarnple query asks for the top 10 products ordered by sales in a given location and time:

```
SELECT    P.pid, P.pnarne, S.sales
FROM      Sales S, Products P
WHERE     S.pid=P.pid AND S.locid==l AND S.tiIneid=3
ORDER BY  S.sales DESC
OPTIMIZE  FOR 10 ROWS
```

The OPTIMIZE FOR N ROWS construct is not in SQL-92 (or even SQL:1999), but it is supported in IBM's DB2 product, and other products (e.g., Oracle 9i) have sirnilar constructs. In the absence of a cue such as OPTIMIZE FOR 10 ROWS, the DBMS computes sales for all products and returns thenl in descending order by sales. The application can close the result cursor (i.e., tenninate the query execution) after consulning 10 rows, but considerable effort has already been expended in cornputing sales for all products and sorting them.

Now let us consider how a DBMS can use the OPTIMIZE FOR cue to execute the query efficiently. The key is to sOlnehow cornpute sales only for products that are likely to be in the top 10 by sales. Suppose that we know the distribution of sales values because we rnaintain a histogran1 on the *sales* cohuun of the Sales relation. We can then choose a value of *sales,* say, c, such that only 10 products have a larger sales value. For those Sales tuples that meet this condition, we can apply the location and tirne conditions as well and sort the result..Evaluating the following query is equivalent to this approach:

```
SELECT    P.pid, P.pnarne, S.sales
FROM      Sales S, Products P
WHERE     S.pid=P.picl AND S.locid=1 AND S.timeid=3 AND S.sales > c
ORDER BY  S.sales DESC
```

This approach is, of course, ruuch faster than the alternative of cornputing all product sales and sorting thern, but there are SOIne in1portant problerns to resolve:

1. *flow do we choose the sales cutoff value* c? EIistograrns and other systeln statistics can be used for this m1rI)()SC, but this can be a tricky issue. For

one thing, the statistics rnaintained by a DBMS are only approximate. For another, even if we choose the cutoff to reflect the top 10 sales values accurately, other conditions in the query Inay elirninate SOHle of the selected tuples, leaving us with fewer than 10 tuples in the result.

2. *What 'if we have more than 10 tuples in the result?* Since the choice of the cutoff c is approxirnate, we could get 1nore than the desired nurnber of tuples in the result. rrhis is easily handled by returning just the top 10 to the user. We still save considerably with respect to the approach of cornputing sales for all products, thanks to the conservative pruning of irrelevant sales infonnation, using the cutoff c.

3. *What 'if we have fewer than 10 tuples in the result?* Even if we choose the sales cutoff c conservatively, we could still cOlnpute fe\ver than 10 result tuples. In this case, we can re-execute the query with a srnaller cutofF value C2 or simply re-execute the original query with no cutoff.

The effectiveness of the approach depends on how well we can estirnate the cutoff and, in particular, on rninimizing the nurnber of tiules we obtain fewer than the desired nurnber of result tuples.

25.5.2 Online Aggregation

Consider the following query, which asks for the average sales arIlount by state:

```
SELECT    L.state, AVG (S.sales)
FROM      Sales S, Locations L
WHERE     S.locid=L.locid
GROUP BY  L.state
```

This can be an expensive query if Sales and Locations are large relations. We cannot achieve fast response tirnes with the traditional approach of cornputing the anwer in its entirety when the query is presented. One alternative, as we have seen, is to use precornputation. Another alternative is to cornpute the answer to the query when the query is presented l)ut return an approximate answer to the user as soon as possible. As the cornputation progresses, the answer quality is continually refined. This approach is called online aggregation. It is very attra,ctive for queries involving aggregation, beca,use efficient techniques for cornputing and refining approxirnate answers are available.

Online aggregation is illustrated in Figure 25.8: For each state--the grouping criterion for our exarnple query · the current value for average sales is displayed, together with a confidence interval The entry for Alaska tells us that the

STATUS	PRIORITIZE	State	AVG(sales)	C..	Interval
		Alabama	5,232.5	97%	103.4
		Alaska	2,832.5	93%	132.2
		Arizona	6,432.5	98%	52.3
		Wyoming	4,243.5		

Figure 25.8 Online Aggregation

current estiInate of average per-store sales in Alaska is $2,832.50, and that this is within the range $2,700.30 to $2,964.70 with 93% probability. rrhe status bar in the first column indicates how close we are to arriving at an exact value for the average sales and the second cohllnn indicates 'whether calculating the average sales for this state is a priority. Estimating average sales for Alaska is not a priority, but estimating it for Arizona is a priority. As the figure indicates, the DBMS devotes Inore systern resources to estiInating the average sales for high-priority states; the estirnate for Arizona is Inucll tighter than that for Alaska and holds with a higher probability. Users can set the priority for a state by clicking on the Prioritize button at any tilne during the execution. This degree of interactivity, together with the continuous feedback provided by the visual display, rnakes online aggregation an attractive technique.

To irnplernent online aggregation, a DEI\!IS lIlust incorporate statistical techniques to provide confidence intervals for approxiInate answers and use nonblocking algorithms for the relational operators. An algorithnl is said to block if it does not produce output tuples until it has consurned all its input tuples. For exarnple, the sort-Illerge join algoritlun blocks because sorting requires all input tuples before detennining the first output tuple. Nested loops join and hash join are therefore preferable to sort-rnerge join for online aggregation. Sirnilarly, hash-based aggregation is better than sort-based aggregation.

25.6 IMPLEMENTATION TECHNIQUES FOR OLAP

In this section we survey 80r11e irnplernentatioll techniques rllotivated by the ()LAP envirornnent. rrhe goal is to provide a feel for how ()LAP systerIls differ fron1 1nore traditional SQL systerns; our discussion is far frorn cornprehensive.

Beyond B+ **Trees:** Complex queries have rnotivated the addition of powerful indexing techniques to DBMSs. In addition to B+ tree indexes, Oracle 9i supports bitlnap and join indexes and Inaintains these dynalnically as the indexed relations are updated. Oracle 9i also supports indexes on expressions over attribute values, such as 10 * *sal* + *bonus*. Microsoft SQL Server uses bitrnap indexes. Sybase IQ supports several kinds of bitrnap indexes, and rnay shortly add support for a linear hashing based index. Informix UDS supports R trees and Inforrnix XPS supports bitlIlap indexes.

The rIlostly-read environruent of OLAP systerns rnakes the CPU overhead of rnaintaining indexes negligible and the requireruent of interactive response tinles for queries over very large datasets rnakes the availability of suitable indexes very important. This combination of factors has led to the developrnent of new indexing techniques. We discuss several of these techniques. We then consider file organizations and other OLAP implenlentation issues briefly.

We note that the ernphasis on query processing and decision support applications in OLAP systems is being cornplemented by a greater erllphasis on evaluating cOlnplex SQL queries in traditional SQL systerIls. Traditional SQL systerns are evolving to support OLAP-style queries more efficiently, supporting constructs (e.g., CUBE and window functions) and incorporating irnpleruentation techniques previously found only in specialized 0 LAP systems.

25.6.1 Bitmap Indexes

Consider a table that describes custorners:

Custoruers(*custid:* integer, *narne:* string, *gender':* boolean, *rating:* integer)

The *rating* value is an integer in the range 1 to 5, and only two values are recorded for *gender*. Cohllnns with few possible values are called sparse. We can exploit sparsity to construct a new kind of index that greatly speeds up queries 011 these cobulins.

The idea is to record values for sparse colurnns as a sequence of bits, one for each possible value. For exarnple, a *gender* value is either 10 or en; a 1 in the first position denotes ruale, and 1 in the second position denotes fe1nale. Similarly, 10000 denotes the *rai'ing* value 1, and 00001 denotes the *rating* value 5.

If we consider the *gender* values for all rows in the Custorners table, we can treat this as a collection of two bit vectors, Olle of which has the associated value M(ale) and the other the associated value F(ernale). Each bit vector has one bit per row in the Custorners table, indicating whether the value in that row is the value associated with the bit vector. The collection of bit vectors for a COhUIIII is called a bitrnap index for that colurnn.

An exaInple instance of the Customers table, together with the bitlnap indexes for *gender* and *rating,* is shown in Figure 25.9.

M	*F*
1	0
1	0
0	1
1	0

custid	*name*	*gender*	*rating*
112	Joe	M	3
115	RaIn	M	5
119	Sue	F	5
112	Woo	M	4

1	*2*	*3*	*4*	*5*
0	0	1	0	0
0	0	0	0	1
0	0	0	0	1
0	0	0	1	0

Figure 25.9 Bitmap Indexes on the Customers Relation

Bitmap indexes offer two important advantages over conventional hash and tree indexes. First, they allow the use of efficient bit operations to answer queries. For example, consider the query, "How Inany Inale custolllers have a rating of 5?" We can take the first bit vector for *gender* and do a bitwise AND with the fifth bit vector for *rating* to obtain a bit vector that has 1 for every male custoIner with rating 5. We can then count the number of Is in this bit vector to answer the query. Second, bitmap indexes can be much luore cOInpact than a traditional B+ tree index and are very amenable to the use of cornpression techniques.

Bit vectors correspond closely to the rid-lists used to represent data entries in Alternative (3) for a traditional B+ tree index (see Section 8.2). In fact, we can think of a bit vector for a given *age* value, say, as an alternative representation of the rid-list for that value.

This suggests away to combine bit vectors (and their advantages of bitwise processing) with B+ tree indexes: We can use Alternative (3) for data entries, using a bit vector representation of rid-lists. A caveat is that, if an rid-list is very slnall, the bit vector representation rnay be Illuch larger than a list of rid values, even if the bit vector is cornpressed. Further, the use of corupression leads to decornprcssion costs, offsetting sorne of the C0I11putational advantages of the bit vector representation.

A Inore flexible approach is to usc a standard list representation of the rid-list for S01ne key values (intuitively, those that contain few clernents) and a bit

vector representation for other key values (those that contain rnany elenlents, and therefore lend themselves to a cOInpact bit vector representation).

This hybrid approach, which can easily be adapted to work with hash indexes as well as B+ tree indexes, has both advantages and disadvantages relative to a standard list of rids approach:

1. It can be applied even to cohllnns that are not sparse; that is, in which are Tnany possible values can appear. The index levels (or the hashing scheIue) allow us to quickly find the 'list' of rids, in a standard list or bit vector representation, for a given key value.
2. Overall, the index is Tnore cornpact because we can use a bit vector representation for long rid lists. We also have the benefits of fast bit vector processIng.
3. On the other hand, the bit vector representation of an rid list relies on a Inapping fron1 a position in the vector to an rid. (This is true of any bit vector representation, not just the hybrid approach.) If the set of rows is static, and we do not worry about inserts and deletes of rows, it is straightforward to ensure this by assigning contiguous rids for rows in a table. If inserts and deletes Inust be supported, additional steps are required. For exanlple, we can continue to assign rids contiguously on a per-table basis and sirnply keep track of which rids correspond to deleted rows. Bit vectors can now be longer than the current nUlnber of rows, and periodic reorganization is required to cOlllpact the 'holes' in the assignrnent of rids.

25.6.2 Join Indexes

Cornputing joins with sIllall response tirnes is extrernely hard for very large relations. One approach to this problern is to create an index designed to speed up specific join queries. Suppose that the Custorners table is to be joined with (1, table called Purchases (recording purchases made by custorners) on the *custid* field.vVe can create a collection of $\langle c, p \rangle$ pairs, where p is the rid of a Purchases record that joins with a Custc)lners recol'c! with *custid* c.

This idea can be generalized to support joins over ruore than two relations. We discuss the special case of a star scherna., in which the fact table is likely to be joined with several dirnension tables. Consider a join query that joins fact table F with dilnension tables D1 and D2 and includes selection conditions on cohunn C_1 of tal)le 1)1 and colurnn (:12 of table D2. We store a tuple $\langle r_1, (\text{'2}, r \rangle$ irl the join index if *T1* is the rid of a tuple in table 1)1 with value (1 in cohunn C_1, '12 is the rid of a tuple in table D2 with value *c2* in colllrnn C_2, and r is the rid of a tllple in the fact table F, and these three tUl)les join with each other.

Complex Queries: The IBM DB2 optimizer recognizes star join queries and perfOfIns rid-based sernijoins (using BIoarn filters) to filter the fact table. Then fact table rows are rejoined to the dimension tables. Cornplex (rnnltitable) dirnension queries (called *snowflake queries*) are supported. DB2 also supports CUBE using smart algorithrns that rninhnize sorts. Microsoft SQL Server optiInizes star join queries extensively. It considers taking the cross-product of srnall dirnension tables before joining with the fact table, the use of join indexes, and rid-based semijoins. Oracle 9i also allows users to create diInensions to declare hierarchies and functional dependencies. It supports the CUBE operator and optirnizes star join queries by elinlinating joins when no colunlll of a dirnension table is part of the query result. DBMS products have also been developed specifically for decision support applications, such as Sybase IQ.

The drawback of a join index is that the nurnber of indexes can grow rapidly if several colurnns in each dirnension table are involved in selections and joins with the fact table. An alternative kind of join index avoids this problem. Consider our exarnple involving fact table F and dirnension tables Dl and D2. Let G_1 be a column of Dl on which a selection is expressed in some query that joins Dl with F. Conceptually, we now join F with Dl to extend the fields of F with the fields of Dl, and index F on the 'virtual field' *G1:* If a tuple of Dl with value CI in colurnn C_1 joins with a tuple of F with rid *r,* we add a tuple $\langle C1, r \rangle$ to the join index. We create one such join index for each colurnn of either Dl or D2 that involves a selection in SOHle join with F; C_1 is an exarnple of such a COIUllUl.

The price paid with respect to the previous version of join indexes is that join indexes created in this way have to be cornbined (rid intersection) to deal with the join queries of interest to us. This can be done efficiently if we rnake the ne\v indexes *bitrnap* indexes; the result is called a, bitrnapped join index. The idea works especially well if cohunns such as C_1 are sparse, and therefore well suited to bitrnap indexing.

25.6.3 File Organizations

Since rllFtny OLAP queries involve just a few colurnns of a large relation, vertical partitioning becornes attractive. However, storing a relation colurnn-\vise can degrade perfoI"rnance for queries that involve several colurnns. An alternative in a rl1ostly-read envirollrnent is to store the relation row-wise, but also store each COIUH1II separately.

A rnore radical file organization is to regard the fact table as a large Illuitidirnensional array and store it and index it as such. This approach is taken in MOLAP systerns. Since the array is lIluch larger than available lnain lnelnory, it is broken up into contiguous chunks, as discussed in Section 23.8. In addition, traditional B+- tree indexes are created to enable quick retrieval of chunks that contain tuples with values in a given range for one or rnore diInensions.

25.7 DATA WAREHOUSING

Data warehouses contain consolidated data from many sources, augrnented with sunnnary inforrnation and covering a long time period. Warehouses are lnuch larger than other kinds of databases; sizes ranging frorn several gigabytes to terabytes are cornman. Typical workloads involve ad hoc, fairly cOlllplex queries and fast response times are important. These characteristics differentiate warehouse applications from OL'TP applications, and different DBMS design and irnplerrlentation techniques nUlst be used to achieve satisfactory results. A distributed DBMS with good scalability and high availability (achieved by storing tables redundantly at more than one site) is required for very large warehouses.

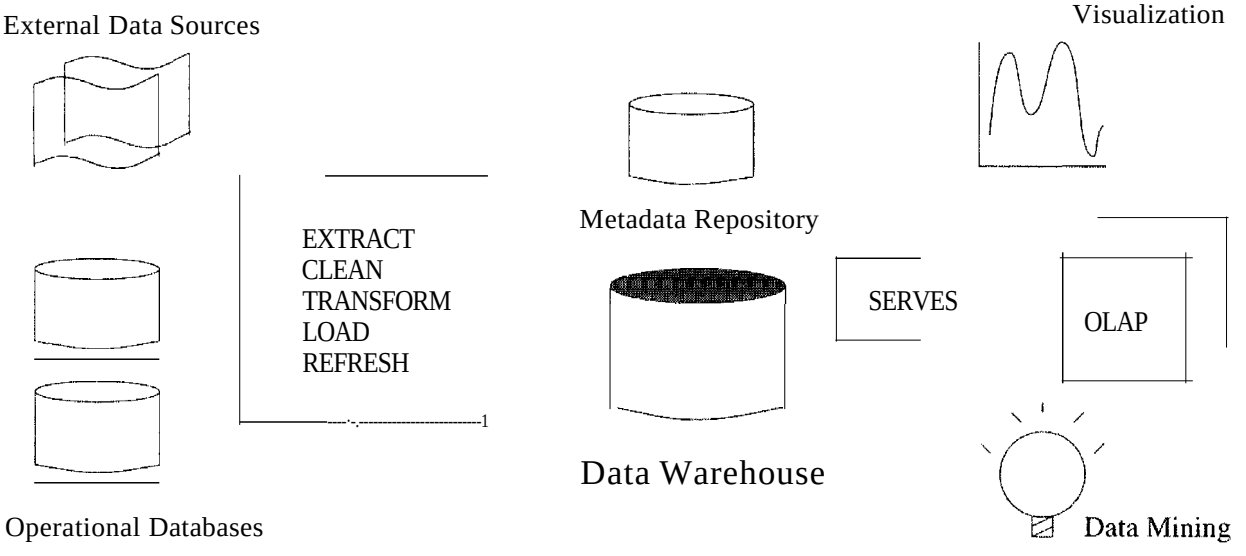


Figure 25.10 A rrypical Data \Varehousing Architecture

A typical data warehousing architecture is illustrated in Figure 25.10. An organization's daily operations access and rnodify operational databases. Data frorl1 these operational databases and other external sources (e.g., custorner profiles supplied by external consultants) are extracted by using interfaces such as JI)BC (see Section 6.2).

25.7.1 Creating and Maintaining a Warehouse

Many challenges rnust be Inet in creating and Inaintaining a large data warehouse.A good database scherua nlust be designed to hold an integrated collection of data copied froIn diverse sources. For exarnple, a cornpany warehouse rnight include the inventory and personnel departrnents' databases, together with sales databases rnaintained by offices in different countries. Since the source databases are often created and rnaintained by different groups, there are a nUlnber of selnantic Inisrnatches across these databases, such as different currency units, different narnes for the saIne attribute, and differences in how tables are nornlalized or structured; these differences Inust be reconciled when data is brought into the warehouse. After the warehouse schenla is designed, the warehouse must be populated, and over tirne, it Inust be kept consistent with the source databases.

Data is extracted from operational databases and external sources, cleaned to Inininlize errors and fill in Inissing information when possible, and **transformed** to reconcile semantic Inismatches. Transforlning data is typically accOlnplished by defining a relational view over the tables in the data sources (the operational databases and other external sources). Loading data consists of ruaterializing such views and storing therll in the warehouse. Unlike a standard view in a relational DBMS, therefore, the view is stored in a database (the warehouse) that is different frorn the database(s) containing the tables it is defined over.

The cleaned and transfonned data is finally loaded into the warehouse. Additional preprocessing such as sorting and generation of surnrnary information is carried out at this stage. Data is partitioned and indexes are built for efficiency. Due to the large vohllue of elata, loading is a slow process. Loading a terabyte of data sequentially can take weeks, and loading even a gigabyte can take hours. Parallelisul is therefore iInportant for loading warehouses.

AJter data is loaded into a warehouse, additional rneasures rnust be taken to ensure that the data in the warehouse is periodically refreshed to reflect updates to the data sources and periodically purge old data (perhaps onto archival rnedia). Observe the connection between the problern of refreshing warehouse tables and a,synchronously rnaintaining replicas of tables in a distributed DBMS. Maintaining replicas of source relations is an essential part of warehousing, and this application clornain is an iInportant factor in the popularity of asynchronous replication (Section 22.11.2), even though asynchronous replication violates the principle of distributed data independence. The problern of refreshing warehouse tables (\vhich are rnaterialized views over tables in

the source databases) has also renewed interest in inerernental Illaintenance of materialized views. (We discuss rnaterialized views in Section 25.8.)

An irnportant task in maintaining a warehouse is keeping track of the data currently stored in it; this bookkeeping is done by storing infofrnation about the warehouse data in the systenl catalogs. The systerIl catalogs associated with a \varehouse are very large and often stored and 111anaged in a separate database called a metadata repository. The size and cornplexity of the catalogs is in part due to the size and cOlnplexity of the warehouse itself and in part because a lot of adrninistrative inforrnation rnust be Inaintained. For example, we HlllSt keep track of the source of each warehouse table and when it was last refreshed, in addition to describing its fields.

1'he value of a warehouse is ultin1ately in the analysis it enables. The data in a warehouse is typically accessed and analyzed using a variety of tools, including OLAP query engines, data. mining algorithrns, inforrnation visualization tools, statistical packages, and report generators.

25.8 VIEWS AND DECISION SUPPORT

Views are widely used in decision support applications. Different groups of analysts within an organization are typically concerned with different aspects of the business, and it is convenient to define views that give each group insight into the business details that concern it. Once a view is defined, we can write queries or new view definitions that use it, as we saw in Section 3.6; in this respect a view is just like a base table. Evaluating queries posed against views is very ilnportant for decision support applications. In this section, we consider how such queries can be evaluated efficiently after placing views within the context of decision support applications.

25.8.1 Views, OLAP, and Warehousing

Views are closely related to OLAP and data warehousing.

OLAP queries are typically aggregate queries. Analysts want fast answers to these queries over very large datasets, and it is natural to consider precoluputing views (see SectiorlS 25.9 and 25.10). In particular, the CUBE operator—discussed in Section 25.3—gives rise to several aggregate queries that are closely related. The relationships that exist between the Inany aggregate queries that arise froln a single CUBE operation can be exploited to develop very effective precornputation strategies. The idea is to choose a subset of the aggregate queries for Inaterialization in such a way that typical CUBE queries can be quickly answered by using the materialized views arld doing S(Hne additional cornplltation. The

choice of views to 111ateria1ize is influenced by how lllany queries they can potentially speed up and by the aillount of space required to store the Inaterialized view (since we have to work with a given alnount of storage space).

A data \varehouse is just a collection of asynchronously replicated tables and periodically synchronized views. A warehouse is characterized by its size, the nuruber of tables involved, and the fact that IllOSt of the underlying tables are froln external, independently lnaintained databases. Nonetheless, the fundaluental probleln in warehouse lnaintenance is asynchronous rnaintenance of replicated tables and materialized views (see Section 25.10).

25.8.2 Queries over Views

Consider the following view, RegionalSales, which cornputes sales of products by category and state:

```
CREATE VIEW RegionalSales (category, sales, state)
        AS SELECT P.category, S.sales, L.state
           FROM   Products P, Sales S, Locations L
           WHERE  P.pid = S.pid AND S.locid = L.locid
```

The following query computes the total sales for each category by state:

```
SELECT   R.category, R.state, SUM (R.sales)
FROM     RegionalSales R
GROUP BY R.category, R,.state
```

While the SQL standard does not specify how to evaluate queries on views, it is useful to think in ternlS of a process called query modification. rrhe idea is to replace the occurrence of RegionalSales in the query by the view definition. The result on this query is

```
SELECT   H,.category, R.state, SUM (R.sales)
FROM     ( SELECT P.category, S.sales, L.state
           FROM   Products P, Sales S, Locations L
           WHERE  P.piel = S.pid AND S.locid = L.locid ) AS R
GROUP BY R.category, R.state
```

25.9 VIEW MATERIALIZATION

We can answer a query on a view by using the query rnodification technique just described. Often, however, queries against cornplex view definitions Illust

be answered very fast because users engaged in decision support activities require interactive response tirlles. Even with sophisticated optimization and evaluation techniques, there is a lirnit to how fast we can answer such queries. Also, if the underlying tables are in a rernote database, the query rIlodification approach rnay not even be feasible because of issues like connectivity and availability.

An alternative to query rnodification is to precornpute the view definition and store the result. When a query is posed on the view, the (unrllodified) query is executed directly on the precornputed result. This approach, called **view materialization**, is likely to be rnuch faster than the query modification approach because the complex view need not be evaluated when the query is computed. Materialized views can be used during query processing in the sarne way as regular relations; for exarnple, we can create indexes on nlaterialized views to further speed up query processing. The drawback, of course, is that we must maintain the consistency of the precomputed (or *m,aterialized)* view whenever the underlying tables are updated.

25.9.1 Issues in View Materialization

Three questions must be considered with regard to view nlaterialization:

1. What views should we rnaterialize and what indexes should we build on the rnaterialized views?
2. Given a query on a view and a set of materialized views, can we exploit the rnaterialized views to answer the query?
3. I-Iow should we synchronize rnaterialized views with changes to the underlying tables? The choice of synchronization technique depends on several factors, such as whether the underlying tables are in a rernote database. We discuss this issue in Section 25.10.

'rhe answers to the first two questions are related. The choice of views to rnaterialize and index is governed by the expected workload, and the discussion of indexing in Chapter 20 is relevant to this question as well. The choice of views to rnaterialize is rnore cornplex than just choosing indexes on a set of database tables, however, because the range of alternative views to rnaterialize is wider. The goal is to rnaterialize a srnaU, carefully chosen set of views that can be utilized to quickly answer rnost of the irnportant queries. COIlversely, once we have chosen a set of views to rnaterialize, we have to consider how they can be used to answer a, given query.

Consider the RegionalSales view. It involves a JOIn of Sales, Products, and Locations and is likely to be expensive to cornpute. On the other hand, if it

is rnaterialized and stored with a clustered B+ tree index on the search key (category, state, sales), we Gall ans\ver the exarnple query by an index-only searl.

Given the rnaterialized view and this index, we can also answer queries of the follo\ving forrn efficiently:

```
SELECT    R.state, SUM (R.sales)
FROM      RegionalSales R
WHERE     R.category == 'Laptop'
GROUP BY  R.state
```

To answer such a query, we can use the index on the Inaterialized view to locate the first index leaf entry with *category* == 'Laptop' and then scan the leaf level until we come to the first entry with *category* not equal to Laptop.

The given index is less effective on the following query, for which we are forced to scan the entire leaf level:

```
SELECT    R.state, SUM (R.sales)
FROM      R,egionalSales R
WHERE     R.state == 'Wisconsin'
GROUP BY  R.category
```

This exanlple indicates how the choice of views to materialize and the indexes to create are affected by the expected workload. This point is illustrated further by our next exarnple.

Consider the following two queries:

```
SELECT    P.category, SUM (S.sales)
FROM      Products P, Sales S
WHERE     P.pic! == S.pic!
GROUP BY  P.category
```

```
SELECT    L.state, SUM (S.sales)
FROM      Locations L,, Sales S
WHERE     L.locid = S.locid
GROUP BY  L.state
```

These two queries require us to join the SaJes table (which is likely to be very large) with another table and aggregate the result. How can we use rnaterialization to speed up these queries? The straightforward approach is to precornpute

each of the joins involved (Products with Sales and Locations with Sales) or to preconlpute each query in its entirety. An alternative approach is to define the following view:

```
CREATE   VIEW rrotalSaJes (pid, locid, total)
         AS SELECT   S.pid, S.locid, SUM (S.sales)
            FROM     Sales S
            GROUP BY S.pid, S.locid
```

The view TotalSales can be rnaterialized and used instead of Sales in our two exalnple queries:

```
SELECT   P.category, SUM (T.total)
FROM     Products P, TotalSales T
WHERE    P.pid = T.pid
GROUP BY P.category

SELECT   L.state, SUM (T.total)
FROM     Locations L, TotalSales T
WHERE    L.locid = rr.locid
GROUP BY L.state
```

25.10 MAINTAINING MATERIALIZED VIEWS

A materialized view is said to be refreshed when we rnake it consistent with changes to its underlying tables. rrhe process of refreshing a view to keep it consistent with changes to the underlying table is often referred to as view maintenance. Two questions to consider are

1. *flow* do we refresh a view' when an underlying table is nlodified? Two issues of particular interest are how to Inaintain views *incrementally*, that is, without recornputing frolI! scratch when there is a change to an underlying table; and how to rnaintain views in a distributed environrnent such as a data warehouse.
2. *When* should we refresh a view in response to a change to an underlying table?

25.10.1 Incremental View Maintenance

A straightforward approach to refreshing a view is to sirnply recompute the view when an underlying table is rnodified. This rnay, in fact, be a reasonable strategy in sorne cases. For exarnple, if the underlying tables are in a

rernote database, the view can be periodically recornputed and sent to the data warehouse \vhere the vie\v is Hlaterialized. This has the advantage that the underlying tables need not be replicated at the warehouse.

\Vhenever possible, however, algorithrns for refreshing a view should be incremental, in that the cost is proportional to the extent of the change rather than the cost of recornputing the view fr(Hn scratch.

To understand the intuition behind incrernental view rnaintenance algorithnls, observe that a given row in the rnaterialized view can appear several times, depending on how often it was derived. (R.ecall that duplicates are not elirninated fro111 the result of an SQL query unless the DISTINCT clause is used. In this section, we discuss rnultiset sernantics, even when relational algebra notation is used.) The rHain idea behind incremental rnaintenance algorithrIls is to efficiently compute changes to the rows of the view, either new rows or changes to the count associated with a row; if the count of a row becornes 0, the row is deleted frorH the view.

We present an incrernental maintenance algorithnl for views defined using projection, binary join, and aggregation; we cover these operations because they illustrate the rHain ideas. The approach can be extended to other operations such as selection, un.ion, intersection, and (rnultiset) difference, as well as expressions containing several operators. The key idea is still to rnaintain the nurnber of derivations for each view row, but the details of how to efficiently conlpute the changes in view rows and associated counts differ.

Projection Views

Consider a view V defined in tenns of a projection on a tableR; that is, $V = n(R)$. Every row v in V has an associated count, corresponding to the nurnber of tirnes it can be derived, which is the nurnber of rows in R that yield v when the projection is applied. Suppose we 1nodifyR by inserting a collection of rows R_i and deleting a collection of existing 1'o\vs R_d.[1] We cornpute $\pi(R_i)$ and add it to V. If the multiset $\pi(R_i)$ contains a row r with count c and r does not appear in V, we add it to V"with count c. If r is in V, we add c to its count. We also cornpute $n(Rd)$ and subtract it fronl V. ()bserve that if r appears in $\pi(R_d)$ with count c, it 111USt also appear in V with a higher count;[2] we subtract c frOTH r's count in V".

1'These collections can be multisets of rows. We can treat a row rnodification as an insert followed by a delete, for sirnplicity.

2As a simple exercise, consider why this rnust be so.

As an exanlple, consider the view $\pi_{sales}(Sales)$ and the instance of Sales shown in Figure 25.2. _Each row in the vie\v has a single cohunn; the (n)\vwith) value 25 appears with count 1, and the value 10 appears with count 3. If we delete one of the rows in Sales \vith *sales* 10, the count of the (row with) value 10 in the view becornes 2. If we insert a new row into Sales with *sales* 99, the view no\v has a row with value 99.

An hnportant point is that we have to rnaintain the counts associated with rows even if the view definition uses the DISTINCT clause, rneaning that duplicates are elilninated frorn the view. Consider the saIne view with set semantics—the DISTINCT clause is used in the SQL view definition··------and suppose that we delete one of the rows in Sales with *sales* 10. Does the view now contain a row with value 10'1 To deterrIline that the answer is yes, we need to maintain the row counts, even though each row (with a nonzero count) is displayed only once in the Inaterialized view.

Join Views

Next, consider a view V defined as a join of two tables, R [x] S. Suppose we modify R by inserting a collection of rows R_i and deleting a collection of rows R_d. We cornpute R_i [x] S and add the result to V. We also C0111pute $R_d \bowtie S$ and subtract the result frorll V. Observe that if r appears in R_d [x] S with count c, it rnust also appear in V with a higher count.[3]

Views with Aggregation

Consider a view V defined over R using GROUP BY on colUllln G and an aggregate operation on colu1nn A. Each row v in the view surnrnarizes a group of tuples in R and is of the fonn $\langle g, summary \rangle$, where g is the value of the grouping colulnn G and the sununary inforInation depends on the aggregate operation. To lnaintain such a view incrernentally, in general, we have to keep a lnore detailed surrllnary than just the inforrnation included in the view. If the aggregate operation is COUNT, we need to Inaintain only a count c for each row v in the view. If a ro\v r is inserted intoR, and there is no row v in V with $v.G = r.G$, we add a new row $(r.G, 1)$. If there is a ro,v v \vith $v.C\} = r.G$, we incrernent its count. If a row r is deleted frolll R, we decrcrnent the count for the row v with $v.Ci = r.G$; v can be deleted if its count becornes 0, because then the last row in this group has been deleted frorn R.

If the aggregate operation is SUM, we have to lllaintain a SUIII s and also a count c. If a row r is inserted into R and there is no row v in V with $v.C; = T.C:$,

[3] As another simple exercise, consider why this mllst be so.

we add a new row $\langle r.G, a, 1\rangle$. If there is a row $\langle r.G, s, c\rangle$, we replace it by $\langle r.G, s + a, c + 1\rangle$. If a row r is deleted frolll R, \ve replace the row $\langle r.G, s, c\rangle$ with $\langle r.G, s - a, c - 1\rangle$; v can be deleted if its count becornes 0. Observe that without the count, we do not know when to delete 'u, since the Slun for a group could be 0 even if the group contains SCHne rows.

If the aggregate operation is AVG, we have to lllaintain a Slun s, a count c, and the average for each row in the view. The SlUll and count are rnaintained incrernentally as already described, and the average is corllputed as s/c.

The aggregate operations MIN and MAX are potentially expensive to rnaintain. Consider MIN. For each group in R, we rnaintain (g, rn, c), where rn is the minimum value for colUllln A in the group g, and c is the count of the nUlllber of rows r in R with $r.G = g$ and $r.A = m$. If a row r is inserted into Rand $r.G = g$, if $r.A$ is greater than the miniriulill m for group g, we can ignore r. If $r.A$ is equal to the 111iniInurll m for r's group, we replace the summary row for the group with $(g, m, c+1)$. If $r.A$ is less than the minirllum m for r's group, we replace the SUlnrnary for the group with $(g, r.A, 1)$. If a row r is deleted frorn Rand $r.A$ is equal to the minimurIl m for T'S group, then we HUlst decrernent the count for the group. If the count is greater than 0, we sinlply replace the surnmary for the group with $(g, rn, c-1)$. However, if the count becomes 0, this Ineans the last row with the recorded rninimum A value has been deleted from R and we have to retrieve the sInallest A value among the relnaining rows in R with-group value r.G-and this might require retrieval of all rows in R with group value $r.G$.

25.10.2 Maintaining Warehouse Views

The views rnaterialized in a data warehouse can be based on source tables in rernote databases. rIhe asynchronous replication techniques discussed in Section 22.11.2 allow us to connnunicate changes at the source to the warehouse, but refreshing views incrernentally in a distributed setting presents sorne unique challenges. To illustrate this, we consider a sirnpleview that identifies suppliers of Toys.

```
CREATE VIEW ToySuppliers (sid)
        AS SELECT S.sid
           FROM   Suppliers S, Products P
           WHERE  S.pid = P.piel AND P.category = 'Toys'
```

Suppliers is a new table introduced for this exarnple; let us assume that it has just two fields, *sid* aIId *pid*, indicating that supplier *sid* supplies part *pid*. The location of the tables Proclucts and Suppliers and the view ToySuppliers

influences how we IIlaintain the view. Suppose that all three are rnaintained at a single site. We can Inaintain the view increlnentally using the techniques discussed in Section 25.10.1. If a replica of the vie\v is created at another site, we can 1l1onitor changes to the Inaterialized vie\v and apply thcIn at the second site using the asynchronous replication techniques froIn Section 22.11.2.

But, what if Products and Suppliers are at one site and the view is Inaterialized (only) at a second site? To rnotivate this scenario, we observe that, if the first site is used for operational data and the second site supports cornplex analysis, the two sites lnay well be adrninistered by different groups. The option of lnaterializing ToySuppliers (a view of interest to the second group) at the first site (run by a different group) is not attractive and may not even be possible; the adnlinistrators of the first site may not want to deal with someone else's views, and the a(hninistrators of the second site n1ay not want to coordinate with sonleone else whenever they Inodify view definitions. As another motivation for rnaterializing views at a different location froIn source tables, observe that Products and Suppliers may be at two different sites. Even if ·we rnaterialize ToySuppliers at one of these sites, one of the two source tables is reillote.

Now that we have presented Inotivation for rnaintaining rroySuppliers at a location (say, Warehouse) different froIn the one (say, Source) that contains Products and Suppliers, let us consider the difficulties posed by data distribution. Suppose that a new Products record (with *category* = 'Toys') is inserted. We could try to rnaintain the view incren1entally as follows:

1. The Warehouse site sends this update to the Source site.
2. To refresh the view, we need to check the Suppliers table to find suppliers of the itern, and so the v\larehouse site asks the Source site for this inforrnation.
3. The Source site returns the set of suppliers for the sold iteln, and the Warehouse site incrernentally refreshes the view.

This works when there are no additional changes at the Source site in between steps (1) and (3). If there are changes, however, the Inaterializecl view can becorne incorrect reflecting a state that can never arise except for anornalies introduced by the preceding, naive, increInental refresh algorithrn. To see this, suppose that Products is enlpty and Suppliers contains just the row $\langle s1, 5\rangle$ initially, and consider the following sequence of events:

1. Product *pid* = 5 is inserted \vith *category* = 'Toys'; Source notifies\Varehouse.
2. Warehouse asks Source for suppliers of product *pid* = 5. *(The only such supplier at this instant is 81.)*

3. The row (82,5) is inserted into Suppliers; Source notifies \Varehouse.

4. To decide whether 82 should be added to the view, we need to know the category of product *pid* = 5, and \Varehouse asks Source. *(Warehouse has not received an answer to its previous question.)*

5. Source now processes the first query frorn Warehouse, finds two suppliers for part 5, and returns this inforrnation to Warehouse.

6. Warehouse gets the answer to its first question: suppliers 81 and 82, and adds these to the view, each with count 1.

7. Source processes the second query frorn \Varehouse and responds with the inforll1ation that part 5 is a toy.

8. Warehouse gets the answer to its second question and accordingly incre-Hlents the count for supplier 82 in the view.

9. Product *pid* = 5 is now deleted; Source notifies Warehouse.

10. Since the deleted part is a toy, Warehouse decrements the counts of nlatch-ing view tuples; 81 has count 0 and is relnoved, but *s2* has count 1 and is retained.

Clearly, 82 should not rernain in the view after part 5 is deleted. This example illustrates the added subtleties of incremental view rnaintenance in a distributed environment, and this is a topic of ongoing research.

25.10.3 When Should We Synchronize Views?

A view maintenance policy is a decision about when a view is refreshed, independent of whether the refresh is incrernental or not. A view can be re-freshed within the sallIe transaction that updates the underlying tables. This is called immediate view Iuaintenance. The update transaction is slowed by the refresh step, and the irupact of refresh increases with the nurnber of materialized views that depend on the updated table.

Alternatively, we can defer refreshing the view. Updates are captured in a log and applied subsequently to the rnaterialized vic\vs. There are several deferred view maintenance policies:

1. Lazy: The rnaterialized view V is refreshed at the tilne a query is evaluated using V, if V is not already consistent with its underlying base tables. This approach slows down queries rather than updates, in contrast to iHnnediate view rnaintenance.

Views for De-cision Support: DBMS vendors are enhancing their main relational products to support decision support queries. IBM DB2 supports materialized views with transaction-consistent or user-invoked maintenance. Microsoft SQL Server supports partition views, \vhich are unions of (ruany) horizontal partitions of a table. These aJ'e airned at a warehollsing envirOllrnent where each partition could be, for exalnple, a rnonthly update. Queries on partition vie\vs are opthnized so that only relevant partitions are accessed. Oracle 9i supports 111aterialized views with transaction-consistent, user-invoked, or tilne-scheduled nlaintenance.

2. Periodic: The lllaterialized view is refreshed periodically, say, once a day. The discussion of the Capture and Apply steps in asynchronous replication (see Section 22.11.2) should be reviewed at this point, since it is very relevant to periodic view lllaintenance. In fact, many vendors are extending their asynchronous replication features to support lllaterialized views. Materialized views that are refreshed periodically are also called snapshots.
3. Forced: rrhe rnaterialized view is refreshed after a certain nurnber of changes have been made to the underlying tables.

In periodic and forced view nlaintenance, queries rllay see an instance of the IIlaterialized view that is not consistent with the current state of the underlying tables. That is, the queries would see a different set of rows if the view definition was recornputed. This is the price paid for fast updates and queries, and the trade-off is sirnilar to the trade-off rnade in using asynchronous replication.

25.11 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are *decision support* applications? :Oiscuss the relationship of *complex SQL queries*, *OLA.P*, *data rnining*, and *data warehousing*. (Section 25.1)
- Describe the rnultidirnensional data luodel. Explain the distinction between *rneasurcs* and *dirnensions* and between *fact tables* and *dimension tables*. What is a *star 8chenz,a?* (Sections 25.2 and 25.2.1)
- Cornrnon OLAP operations have received special naInes: *roll-up*, *drill-down*, *pivoting*, *slicing*, and *dicing*. Describe each of these operations and illustrate thern using exarnples. (Section 25.3)
- I)escribe the SQL:1999 ROLLUP and CUBE features and their relationship to the ()LAP operations. (Section 25.3.1)

- Describe the SQL:1999 WINDOW feature, in particular, frarning and ordering of windows. How does it support queries over ordered data? Give exarnples of queries that are hard to express without this feature. (Section 25.4)
- New query paradigrns include *top N queries* and *online aggTegation.* Explain the nlotivation behind these concepts and illustrate then1 through exaruples. (Section 25.5)
- Index structures that are especially suitable for OLAP systen1s include *bitrnap indexes* and *join indexes.* Describe these structures. How are bitrnap indexes related to B+ trees? (Section 25.6)
- Information about daily operations of an organization is stored in *opeTational databases.* Why is a *data warehouse* used to store data frolH operational databases? What issues arise in data warehousing? Discuss *data extTaction, cleaning, transjoTrnation,* and *loading.* Discuss the challenges in efficiently *TejTeshing* and *purging* data. (Section 25.7)
- Why are views irnportant in decision support environments? How are views related to data warehousing and OLAP? Explain the *queTy modification* technique for answering queries over views and discuss why this is not adequate in decision support environrnents. (Section 25.8)
- What are the rnain issues to consider in maintaining materialized views? Discuss how to select views to materialize and how to use rnaterialized views to answer a query. (Section 25.9)
- How can views be rnaintained *incTernentally?* Discuss all the relational algebra operators and aggregation. (Section 25.10.1)
- Use an exarnple to illustrate the added cornplications for incrernental view maintenance introduced by data distribution. (Section 25.10.2)
- Discuss the choice of an appropriate *rnaintenance policy* for when to refresh a view. (Section 25.10.3)

EXERCISES

Exercise 25.1 Briefly answer the following questions:

1. How do warehousing, OLAP, and data rnining cornplernent each other?
2. What is the relationship between data warehousing and data replication? Which fornl of replication (synchronous or asynchronous) is better suited for data warehousing? Why?
3. What is the role of the rnetadata repository in a data warehouse'? How does it differ frorn a catalog in a relational DBMS?
4. What considenttions are involved in designing a data warehouse'?

5. Once a warehouse is designed and loaded, how is it kept current with respect to changes to the source databases?
6. One of the advantages of a warehouse is that we can use it to track how the contents of a relation change over titue; in contrast, we have only the current snapshot of a relation in a regular DBMS. Discuss how you would maintain the history of a relation *R,* taking into account that 'old' infonnation lllust sOlnehow be purged to rnake space for Hew infonnatioll.
7. Describe dilnensions and rneasures in the multidirnensional data model.
8. What is a fact table, and why is it so irnportant frOIn a performance standpoint?
9. What is the fundarnental difference between MOLAP and ROLAP systems?
10. \Vhat is a star scheIna? Is it typicaU:y in BCNF? Why or why not?
11. How is data rnining different from OLAP?

Exercise 25.2 Consider the instance of the Sales relation shown in Figure 25.2.

1. Show the result of pivoting the relation on *pid* and *tirneid.*
2. Write a collection of SQL queries to obtain the same result as in the previous part.
3. Show the result of pivoting the relation on *pid* and *lacid.*

Exercise 25.3 Consider the cross-tabulation of the Sales relation shown in Figure 25.5.

1. Show the result of roll-up on *lacid* (i.e., state).
2. Write a collection of SQL queries to obtain the same result as in the previous part.
3. Show the result of roll-up on *lacid* followed by drill-down on *pid.*
4. Write a collection of SQL queries to obtain the same result as In the previous part, starting with the cross-tabulation shown in Figure 25.5.

Exercise 25.4 Briefly answer the following questions:

1. What is the differences between the WINDOW clause and the GROUP BY clause'?
2. Give an example query that cannot be expressed in SQL without the WINDOW clause but that can be expressed with the WINDOW clause.
3. What is the *frame* of a window in SQL:19997
4. Consider the fonowing simple GROUP BY query.

```
SELECT    T.year, SUM (S.sales)
FROM      Sales 5, Times T
WHERE     S.tilneid='T.timeid
GROUP BY  T.year
```

Can you write this query in SQL:1999 without using a GROUP BY cIa.use? (Hint: Use the SQL:1999 WINDOW clause.)

Exercise 25.5 Consider the Locations, Products, and Sales relations shown in Figure 25.2. Write the following queries in SQL:1999 llsing the WINDOW clause whenever you need it.

1. Find the percentage change in the total IJ10nthly sales for each location.
2. Find the percentage change in the total quarterly sales for each product.

3. Find the average daily sales over the preceding 30 days for each product.
4. For each week, find the maximum uloving average of sales over the preceding four \veeks.
5. Find the top three locations ranked by total sales.
6. F'ind the top three locations ranked by curnulative sales, for every month over the past year.
7. Rank all locations by total sales over the past year, and for each location print the difference in total sales relative to the location behind it.

Exercise 25.6 Consider the CustOIuers relation and the bitmap indexes shown in Figure 25.9.

1. For the same data, if the underlying set of rating values is assuIued to range frolll 1 to 10, show how the bitnlap indexes would change.
2. How would you use the bitIllap indexes to answer the following queries? If the bitmap indexes are not useful, explain why.
   (a) How many customers with a rating less than 3 are male?
   (b) What percentage of custoIners are male?
   (c) How rnany customers are there?
   (d) How many custonlers are named Woo?
   (e) Find the rating value with the greatest number of customers and also find the nUlll-bel' of custorners with that rating value; if several rating values have the maxirnurn number of custoIllers, list the requested infonuation for all of theIn. (AssuIne that very few rating values have the same nUluber of customers.)

Exercise **25.7** In addition to the Customers table of Figure 25.9 with bitrnap indexes on *gender* and *'rating,* assurne that you have a table called Prospects, with fields *rating* and *prospectid.* This table is used to identify potential customers.

1. Suppose that you also have a bitrnap index on the *rating* field of Prospects. Discuss whether or not the bitnlap indexes would help in corllputing the join of Custorners and Prospects on *rating.*
2. Suppose you have *no* bitrnap index on the *rating* field of Prospects. Discuss whether or not the bitrnap indexes on CustOIuers would help in conlputing the join of Custorners and Prospects on *rating.*
3. Describe the use of a join index to support the join of these two relations with the join condition *custid=prospectid.*

Exercise 25.8 Consider the instances of the Locations, Products, and Sales relations shown in Figure 25.2.

1. Consider the basic join indexes described in Section 25.6.2. Suppose you want to optiInize for the following two kinds of queries: Query 1 finds sa.les in a given city, and Query 2 finds sa.les in a given state. Show the indexes you would create on the example instances shown in Figure 25.2.
2. Consider the bitmapped join indexes described in Section 25.6.2. Suppose you want to optirnize for the following two kinds of queries: Query 1 finds sales in a given city, and Query 2 finds sales in a given state. Show the indexes that you would create on the exanlple instances shown in Figure 25.2.

3. Consider the basic join indexes described in Section 25.6.2. Suppose you want to optimize for these two kinds of queries: Query 1 finds sales in a given city for a given product name, and Query 2 finds sales in a given state for a given product category. Show the indexes that you would create on the exarl1ple instances shown in Figure 25.2.
4. Consider the bitmapped join indexes described in Section 25.6.2. Suppose you want to optirnize for these two kincls of queries: Query 1 finds sales in a given city for a given product narne, and Query 2 finds sales in a given state for a given product category. Show the indexes that you would create on the example instances shown in Figure 25.2.

Exercise 25.9 Consicler the view NurnReservations defined as:

```
CREATE VIEW NumReservations (sid, snarnc, nUlures)
      AS SELECT S.sid, S.snarne, COUNT (*)
         FROM    Sailors S, Reserves R
         WHERE   S.sid = R.sid
         GROUP BY 8.sid, S.sname
```

1. How is the following query, which is intended to find the highest number of reservations nlade by smne one sailor, rewritten using query modification?

```
SELECT   MAX (N.numres)
FROM     NurnReservations N
```

2. Consider the alternatives of cornputing on deluand and view materialization for the preceding query. Discuss the pros and cons of materialization.
3. Discuss the pros and cons of materialization for the following query:

```
SELECT   N.snarlle, MAX (N.numres)
FROM     NumReservations N
GROUP BY N.sname
```

Exercise 25.10 Consider the Locations, Products, and Sales relations in Figure 25.2.

1. To decide whether to rnaterialize a view, what factors do we need to consider?
2. Assurne that we have defined the following lnaterialized view:

```
SELECT   L.state, S.sales
FROM     Locations L, Sales S
WHERE    8.locid=L.locid
```

   (a) Describe what auxiliary infornlatioll the algorithnl for incrernental view rnainte-nance frorn Section 25.10.1 maintains and how this data helps in lnainta.ining the view incrernentally.

   (b) Discuss the pros and cons of ruaterializing this view.

3. Consider the materialized view in the previous question. Assume that the relations Locations and Sales are stored at Olle site, but the view is rnaterialized on a second site. Why would we ever want to luaintain the view at a second site? Give a concrete exarnple where the view could become inconsistent.
4. ASSUITW that we have defined the following rnaterialized view:

```
SELECT   T.year, I..state, SUM (S.sales)
FROM     Sales 8, 'rirnes '1', Locations L
WHERE    S.tirneid=T.tilneid AND S.locid=L.locid
GROUP BY rr.year, L.state
```

(a) Describe what auxiliary infoflnation the algorithnl for incrernental view rnaintenance frOIn Section 25.10.1 luaintains, and how this data helps in rnaintaining the view increluentaJly.

(b) Discuss the pros and cons of 11laterializing this view.

BIBLIOGRAPHIC NOTES

A good survey of data warehousing and OLAP is presented in [161], which is the source of Figure 25.10. [686] provides an overview of OLAP and statistical database research, showing the strong parallels between concepts and research in these two areas. The book by Kirnball [436], one of the pioneers in warehousing, and the collection of papers in [(2) offer a good practical introduction to the area. The term OLAP was popularized by Codd's paper [191]. For a recent discussion of the performance of algorithms utilizing bitmap and other nontraditional index structures, see [575].

Stonebraker discusses how queries on views can be converted to queries on the underlying tables through query modification [713]. Hanson cmnpares the perfornlance of query modification versus immediate and deferred view maintenance [365]. Srivastava and Roterll present an analytical model of materialized view maintenance algorithnls [707]. A number of papers discuss how rnaterialized views can be incrementally maintained as the underlying relations are changed. Research into this area has become very active recently, in part because of the interest in *data warehouses,* which can be thought of as collections of views over relations from various sources. An excellent overview of the state of the art can be found in [348], which contains a number of influential papers together with additional rnaterial that provides context and background. The following partial list should provide pointers for further reading: [100, 192, 193, 349, 369, 570, 601, 635, 664, 705, 800].

Gray et al. introduced the CUBE operator [335], and optirnization of CUBE queries and efficient maintenance of the result of a CUBE query have been addressed in several papers, including [12, 94, 216, 367, 380, 451, 634, 638, 687, 799]. Related algorithrns for processing queries with aggregates and grouping are presented in [160, 166]. Rao, Badia, and Van Gucht address the irnplelnentation of queries involving generalized quantifiers such as *a majority of* [618]. Srivastava, Tan, and LUIII describe an access ruethod to support processing of aggregate queries [708]. Shannlugasundaranl et al. discuss how to ruaintain cornpressed cubes for approxirnate answering of aggregate queries in [675].

SQL:1999's support for OLAP, including CUBE and WINDOW constructs, is described in [523]. The windowing extensions are very sirnilar to SQL extension for querying sequence data, called SRQL, proposed in [610]. Sequence queries have received a lot of attention recently. Extending relational systeills, which deal with sets of records, to deal with sequences of records is investigated in [473, 665, 671].

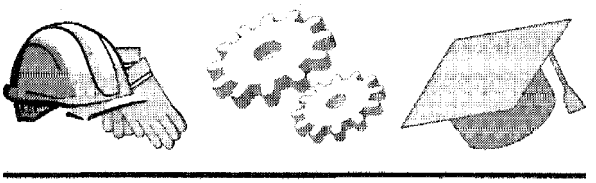
There has been recent interest in one-pass query evaluation algorithnls and database rnanagernent for data streaIns. A recent survey of data rnanagernent for data streams and algorithrns for data stream processing can be fonnd in [49J. Exarnples include quantile and order-statistics cOlnputation [340, 50G], estirnating frequency rnornents and join sizes [34, 35], estirnating correlated aggregates [310], rllultidirnensionaJ regression analysis [173], and cornputing one-dirnensional (i.e., single-attribute) histograrns and Haar wavelet clecmnpositioI1s [319, 345].

Other work includes techniques for incrementally IllElintaining equi-depth histograms [313] and Baal' wavelets [515], rnaintaining sarnples and siluplc statistics over sliding \vindows [201],

as well as general, high-level architectures for stream database systenlS [50]. Zdonik et al. describe the architecture of a database system for Hl0nit;oring data streaU1S [795]. A language infrastructure for developing data streaIll applications is described by Cortes et al. [199].

Carey and Kossrnann discuss how to evaluate queries for which only the first few answers are desired [135, 1:36]. Donjerkovic and Ralnakrishnan consider how a probabilistic approach to query optiInization call be applied to this probleul [229]. [120] compares several strategies for evaluating Top N queries. Hellerstein et al. discuss how to return approxiInate answers to aggregate queries and to refine thern 'online.' [47, 374]. This work has been extended to online cOlnputation of joins [354], online reordering [617] and to adaptive query processing [48].

There has been recent interest in approximate query answering, where a small synopsis data structure is used to give fast approxiruate query answers with provable perforrnance guarantees [7, 8, 61, 159, 167, 314, 759].



26

DATA MINING

- What is data mining?
- What is lliarket basket analysis? What algorithms are efficient for counting co-occurrences?
- What is the a priori property and why is it important?
- What is a Bayesian network?
- What is a classification rule? What is a regression rule?
- What is a decision tree? How are decision trees constructed?
- What is clustering? What is a salllple clustering algorithln?
- What is a similarity search over sequences? How is it implmuented?
- How can data mining models be constructed increluentally?
- What are the new mining challenges presented by data streallls?
- Key concepts: data nlining, KDD process; market basket analysis, co-occurrence counting, association rule, generalized association rule; decision tree, classification tree; clustering; sequence similarity search; incrernental model maintenance, data streanls, block evolution

1' he secret of success is to know sornething nobody else knows.

—Aristotle Onassis

Data luining consists of finding interesting trends or patterns in large datasets to guicle decisions about future activities. There is a genera] expectation that

data ruining tools should be able to identify these patterns in the data with minirnal user input. The patterns identified by such tools can give a data analyst useful and unexpected insight that can be Illore carefully investigated subsequently, perhaps using other decision support tools. In this chapter, we discuss several widely studied data luining tasks. COllunercial tools are available for each of these tasks frorll major vendors, and the area is rapidly gTowing in ilnportance as these tools gain acceptance in the user cornrnunity.

We start in Section 26.1 by giving a short introduction to data mining. In Section 26.2, we discuss the irnportant task of counting co-occurring items. In Section 26.3, we discuss how this task arises in data mining algorithms that discover rules froln the data. In Section 26.4, we discuss patterns that represent rules in the forln of a tree. In Section 26.5, we introduce a different data rnining task, called *clustering*, and describe how to find clusters in large datasets. In Section 26.6, we describe how to perform siInilarity search over sequences. We discuss the challenges in rnining evolving data and data streams in Section 26.7. We conclude with a short overview of other data mining tasks in Section 26.8.

26.1 INTRODUCTION TO DATA MINING

Data nlining is related to the subarea of statistics called *exploratory data analysis,* which has siruilar goals and relies on statisticalrueasures. It is also closely related to the subareas of artificial intelligence called *knowledge discovery* and *rnachine learning.* The important distinguishing characteristic of data rnining is that the volume of data is very large; although ideas froln these related areas of study are applicable to data nlining problems, *scalability with respect to data size* is an important new criterion. An algorithm is scalable if the running tirne grows (linearly) in proportion to the dataset size, holding the available systenl resources (e.g., arnount of rnain rnemory and CPU processing speed) constant. Old algorithms must be adapted or new algorithnls developed to ensure scalability when discovering patterns from data.

Finding useful trends in datasets is a rather loose definition of data 111ining: In a certain sense, all database queries can be thought of as doing just this. Indeed, we have a continuurn of ana.lysis and exploration tools with SQL queries at one end, OLAP queries in the rniddle, and data ruining techniques at the other end. SQL queries are constructed! using relational algebra (with sorne extensions), OLAP provides higher-level querying idiorlls based on the rnultidirnensional data rn.odel, and data mining provides the rnost abstract analysis operations. We can think of different data rnining tasks as cornplex 'queries' specified at a high level, with a few parameters that are user-defina.ble, and for which specialized algorithrns are implemented.

SQL/MM: Data Mining SQL/MM: The SQL/MM: Data Mining extension of the SQL:1999 standard supports four kinds of data mining nlodels: *frequent itemsets and association rules, clusters of records, reg'ression trees*, and *classification trees*. Several new data types are introduced. These data types play several roles. SaIne represent a particular class of model (e.g., `DM_RegressionModel`, `DM_ClusteringModel`); some specify the input parameters for a mining algorithm (e.g., `DM_RegTask`, DM_ClusTask); some describe the input data (e.g., `DM_LogicalDataSpec`, `DM_MiningData`); and sornerepresent the result of executing a rnining algorithm (e.g., `DM_RegResult`, DM_ClusResult). Taken together, these classes and their methods provide a standard interface to data mining algorithms that can be invoked frorn any SQL:1999 database systern. The data mining rnodels can be exported in a standard XML format called Predictive Model Markup **Language** (PMML); models represented using PMML can be hnported as well.

In the real world, data rnining is much more than sirnply applying one of these algorithnls. Data is often noisy or inconlplete, and unless this is understood and corrected for, it is likely that rnany interesting patterns will be rnissed and the reliability of detected patterns will be low. Further, the analyst nlust decide what kinds of rnining algoritlulls are called for, apply them to a well-chosen subset of data sarnples and variables (i.e., tuples and attributes), digest the results, apply other decision support and mining tools, and iterate the process.

26.1.1 The Knowledge Discovery Process

The knowledge discovery and data mining (KDD) process can roughly be separated into four steps.

1. Data Selection: The target subset of data and the attributes of interest are identified by exalnining the entire raw dataset.
2. Data Cleaning: Noise and outliers are relnoved, field values are transfonned to cornrnon units and SOUIC new fields are created by cornbining existing fields to facilitate analysis. The data is typically put into a, relational fonnat, and several tables rnight be cornbined in a *denormalization* step.
3. Data Mining: We apply data rnining algorit1ll11S to extract interesting patterns.
4. Evaluation: The patterns are presented to end-users in an understandable fonn, for example, through visualization.

The results of any step in the KDD process lllight lead us back to an earlier step to redo the process with the new knowledge gained. In this chapter, however, we limit ourselves to looking at algoritlnns for SaIne specific data rnining tasks. We do not discuss other aspects of the I(DD process.

26.2 COUNTING CO-OCCURRENCES

\Ve begin by considering the probleln of counting co-occurring iterns, which is rnotivated by problelTIs such as lllarket basket analysis. A market basket is a collection of items purchased by a custOlner in a single customer transaction. A cnstorner transaction consists of a single visit to a store, a single order through a mail-order catalog, or an order at a store on the Web. (In this chapter, we often abbreviate *customer transaction* to *transaction* when there is no confusion with the usual nleaning of *transaction* in a DBMS context, which is an execution of a user program.) A COIIIIllon goal for retailers is to identify items that are purchased together. This inforrnation can be used to improve the layout of goods in a store or the layout of catalog pages.

transid	*custid*	*date*	*item*	*qty*
111	201	5/1/99	pen	2
111	201	5/1/99	ink	1
111	201	5/1/99	milk	3
111	201	5/1/99	juice	6
112	105	6/3/99	pen	1
112	105	6/3/99	ink	1
112	105	6/3/99	milk	1
113	106	5/10/99	pen	1
113	106	5/io/99	Inilk	1
114	201	6/1/99	pen	2
114	201	6/1/99	ink	2
114	201	6/1/99	juice	4
114	201	6/1/99	water	1

Figure 26.1 The Purchases Relation

26.2.1 Frequent Itemsets

We use the Purchases relation shown in Figure 26.1 to illustrate frequent itemsets. The records are shown sorted into groups by transaction. All tuples in a group have the saIne *transid*, and together they describe a custorner transaction, which involves purchases of one or Inore iterns. A transaction occurs

on a given date, and the nanle of each purchased itenl is recorded, along with the purchased quantity. ()bserve that there is redundancy in Purchases: It can be decolnposed by storing *transid–custid–date* triples in a separate table and dropping *custid* and *date* froln Purchases; this nlay be how the data is actually stored. However, it is convenient to consider the Purchases relation, as shown in Figure 26.1, to corupute frequent iternsets. Creating such 'denormalized' tables for ease of data rnining is cOIIllllonly done in the data cleaning step of the I(DD process.

By examining the set of transaction groups in Purchases, we can rnake observations of the fornl: "In 75% of the transactions a pen and ink are purchased together." rrhis stateulent describes the transactions in the database. Extrapolation to future transactions should be done with caution, as discussed in Section 26.3.6. Let us begin by introducing the terminology of rnarket basket analysis. An itemset is a set of itelTIS. The support of an itelnset is the fraction of transactions in the database that contain all the iterus in the iterllset. In our exaIupl.e, the itelllset {pen, ink} has 75% support in Purchases. We can therefore conclude that pens and ink are frequently purchased together. If we consider the itelllset {milk, juice}, its support is only 25%; milk and juice are not purchased together frequently.

Usually the nUlnber of sets of itenlS frequently purchased together is relatively sInall, especially as the size of the itenlsets increases. We are interested in all iterllsets whose support is higher than a user-specified minimUIl1 support called *minsup*; we call such itemsets frequent itemsets. For exarnple, if the Illinirl1Unl support is set to 70%, then the frequent iterllsets in our example are {pen}, {ink}, {nlilk}, {pen, ink}, and {pen, 111ilk}. Note that we are also interested in iternsets that contain only a single iteru since they identify frequently purchased iterl1s.

We show an algorithrn for identifying frequent iterllsets in Figure 26.2. This algorithrn relies on a sirnple yet fundarnentaJ property of frequent iterIlsets:

> The a Priori Property: Every subset of a frequent iterllset is also a frequent itelnset.

'fhe algorithnl proceeds iteratively, first identifying frequent iterIlsets 'with just one itcrll. In each subsequent iteration, frequent iterl1sets identified in the previous iteration are extended with another itern to generate larger candidate itcrnsets. By considering only iterllsets obtained by enlarging frequent iternsets, we greatly reduce the nurnber of candidate frequent itcrllsets; this optirnization is crucial for efficient execution. The a priori property guarantees that this optilnizatic)ll is correct; that is, we do not Iniss any frequent iterllsets. A single scan of all transactions (the Purchases relation in our example) suffices to

```
foreach itelll,                                              Level 1
    check if it is a frequent iternset // appears in > minsup transactions
k=l
repeat          // Iterative, level-wise identification of frequent itelllsets
    foreach new frequent iterllset Ik with k iterlls          // Level k + 1
        generate all iterl1sets lk+l with k + 1 itelllS, lk C Ik+l
    Scan all transactions once and check if
    the generated k + 1-iterIlsets are frequent
    k=k+1
until no new frequent itemsets are identified
```

Figure 26.2 An Algorithm for Finding Frequent Itemsets

determine which candidate iterllsets generated in an iteration are frequent. The algorithm terminates when no new frequent itemsets are identified in an iteration.

'We illustrate the algorithrn on the Purchases relation in Figure 26.1, with *minsup* set to 70%. In the first iteration (Levell), we scan the Purchases relation and deterllline that each of these one-iterll sets is a frequent iternset: *{pen}* (appears in all four transactions), *{ink}* (appears in three out of four transactions), and *{rnilk}* (appears in three out of four transactions).

In the second iteration (Level 2), we extend each frequent itemset with an additional itenl and generate the following candidate iterIlsets: {*pen, ink}, {pen, milk}, {pen, juice}, {ink, rnilk}, {ink, juice},* and *{rnilk, juice}.* By scanning the Purchases relation again, we deterrnine that the following are frequent ite111sets: {*pen, ink}* (appears in three out of four transactions), and *{pen, rnilk}* (appears in three out of four transactions).

In the third iteration (Level 3), we extend these itelllsets with an additional iteHl and generate the following candidate itcrl1sets: {*pen, ink, milk*}, {*pen, ink, juice},* and *{pen, milk, juice}*. (Observe that {*ink, milk, juice}* is not generated.) A third sca.n of the Purchases relation aJlows us to deterrnine that none of these is a frequent iterTlset.

The sirnple algoritlnll presented here for finding frequent iternsets illustrates the principal feature of Inore sophisticated algorithrns, naruely, the iterative generation and testing of candidate itcrnsets. We consider one irnportant refincrnent of this sirnple algorithrn. Cjenerating candidate iternsets by adding an itCHl to a known frequent iternset is an atterIlpt to lirnit the rnunber of candidate itcrIlsets using the a priori property. rrhe a priori property implies that a can-

didate iternset can be frequent only if all its subsets are frequent. Thus, we can reduce the nUlnber of candidate iternsets further---*a priori*, or before scanning the Purchases database---by checking whether all subsets of a newly generated candidate itcIIlset are frequent. Only if all subsets of a candidate iternset are frequent do we cOlnpute its support in the subsequent database scan. COlnpared to the sirnple algoritlun, this refined algoritlull generates fewer candidate itenlsets at each level and thus reduces the arnount of conlputation perfonned during the database scan of Purchases.

Consider the refined algorithrn on the Purchases table in Figure 26.1 with $minsup$= 70%. In the first iteration (Level 1), we deterrnine the frequent itemsets of size one: *{pen}, {ink},* and $\{milk\}$. In the second iteration (Level 2), only the following candidate itemsets rernain when scanning the Purchases table: $\{pen, ink\}$, $\{pen, milk\}$, and *{ink, rnilk}.* Since *{juice}* is not frequent, the iterllsets *{pen, juice}, {ink, juice},* and *{rnilk, juice}* cannot be frequent as well and we can elirninate those iterIlsets a priori, that is, without considering therll during the subsequent scan of the Purchases relation. In the third iteration (Level 3), no further candidate itemsets are generated. The iternset *{pen, ink,* $milk\}$ cannot be frequent since its subset *{ink, milk}* is not frequent. Thus, the irnproved version of the algorithrll does not need a third scan of Purchases.

26.2.2 Iceberg Queries

We introduce iceberg queries through an exaillple. Consider again the Purchases relation shown in Figure 26.1. Assurne that we want to find pairs of custorners and iterns such that the custorner has purchased the item rllore than five thnes. We can express this query in SQL as follows:

```
SELECT    P.custid, P.itern, SUM (P.qty)
FROM      Purchases P
GROUP BY  P.custid, P.itern
HAVING    SUM (P.qty) > 5
```

rrhink about how this query would be evaluated by a relational DBMS. Conceptually, for each $(custid, item)$ pair, we need to check whether the surn of the qty field is greater than 5. One approach is to rnake a scan over the Purchases relation and rnaintain running surns for each *(c'Ustid, itern)* pair. T'his is a feasible execution strategy as long as the nurnber of pairs is sruaU enough to fit into lIlain rncIIlory. If the nurnber of pairs is larger than rnain Inernory, lnorc expensive query evaluation plans,\vhich involve either sorting or hashing, have to be used.

The query has an irnporta"nt property not exploited by the preceding execution strategy: Even though the Purchases relation is potentially very large and the

nurnber of (*custid, item*) groups CaJl be huge, the Cyutput of the query is likely to be relatively sInall because of the condition in the HAVING clause. ()nly groups where the custorner has purchased the itCHl Inure than five tiInes appear in the output. For exarllple, there are nine groups in the query over the Purchases relation shown in Figure 26.1, although the output contains only three records. The nurnber of groups is very large, but the answer to the query---·-the tip of the iceberg—is usually very sInan. Therefore, we call such a query an iceberg query. In general, given a relational scherna H. with attributes *A.1. A2,* ... , *Ak,* and *B* and an aggrega,tion function aggr, an iceberg query has the follo\ving structure:

```
SELECT    R.Al, R.A2, ..., R,.Ak, aggr(R.B)
FROM      H,elation H,
GROUP BY  R,.AI, ..., R.Ak
HAVING    aggr(R.B) >= constant
```

Traditional query plans for this query that use sorting or hashing first cornpute the value of the aggregation function for all groups and then elirninate groups that do not satisfy the condition in the HAVING clause.

Cornparing the query with the probleur of finding frequent itenlsets discussed in the previous section, there is a striking sirnilarity. Consider again the Purchases relation shown in Figure 26.1 and the iceberg query froIn the beginning of this section. We are interested in (*custid, itern)* pairs that have SUM (P.qty) > 5. lJsing a variation of the a priori property, we can argue that we only have to consider values of the *c'Ust'id* field where the custorner has purchased at least five it-eurs. We can generate such iterns through the following query:

```
SELECT    P.cllstid
FROM      Purchases P
GROUP BY  P.cllstid
HAVING    SUM (P.qty) > 5
```

Sirnilarly, we can restrict the candidate values for theitern field through the following query:

```
SELECT    P.itern
FROM      Purchases P
GROUP BY  P.iteul
HAVING    SUM (P.qty) > 5
```

If we restrict the corrlputation of the original iceberg query to (*custid, 'itern)* groups where the field values are in the output of the previous t\vo queries, we elirninate a large nUlllber of (*custid, item*) pairs a priori. So, a possible

evaluation strategy is to first COlnpute candidate values for the *custid* and *item* fields, and use eornbinations of only these values in the evaluation of the original iceberg query. We first generate candidate field values for individual fields and use only those values that survive the a priori pruning step as expressed in the two previous queries. Thus, the iceberg query is arnenable to the salIle bottorn-up evaluation strategy used to find frequent iternsets. In particular, we can use the a priori property as follows: We keep a counter for a group only if each individual cOlnponent of the group satisfies the condition expressed in the HAVING clause. The perfonnance irnprovernents of this alternative evaluation strategy over traditional query plans can be very significant in practice.

Even though the botto111-UP query processing strategy elinlinates lnany groups a priori, the nlunber of (*custid*, *itern)* pairs can still be very large in practice; even larger than Inain lllernory. Efficient strategies that use sall1pling and lllore sophisticated hashing techniques have been developed; the bibliographic notes at the end of the chapter provide pointers to the relevant literature.

26.3 MINING FOR RULES

Many algorithrIls have been proposed for discovering various fonns of rules that succinctly describe the data. We now look at some widely discussed fonns of rules and algorithnls for discovering thenl.

26.3.1 Association Rules

We use the Purchases relation shown in Figure 26.1 to illustrate association rules. By examining the set of transactions in Purchases, we can identify rules of the forrn:

$$\{pen\} \Rightarrow \{ink\}$$

This rule should be read as follows: "If a pen is pUfcha.sed in a transaction, it is likely that ink is also be purchased in that transaction." It is a staternent that describes the transactions in the database; extrapolation to future transactions should be done with caution, as discussed in Section 26.3.6. More generally, an association rule has the forIn $LHS \Rightarrow RHS$, where both *LIIS* and *RHS* are sets of iterns. The interpretation of such a, rule is that if every itern in *LIIS* is purchased in a transaction, then it is likely that the iterIlS in *RHS* are purchased as well.

rThere are two important measures for (1,n association rule:

- Support: The support for a set of iterns is the percentage of transa,ctions that contain all these iterIls. The support for a rule $LIIS \Rightarrow RHS$ is the

support for the set of itenlS *LHS* *Rf!S*. For exalnple, consider the rule *{pen}* ⇒ *{ink}*. The support of this rule is the support of the itenlset *{pen, ink},* which is 75%.

- Confidence: Consider transactions that contain all iterIls in *LHS*. The confidence for a rule *LlIS* ⇒ *RHS* is the percentage of such transactions that also contain all iterIls in *RHS*. More precisely, let *sup(LHS)* be the percentage of transactions that contain *LllS* and let *s'up(LliS* ∪ *RHS)* be the percentage of transactions that contain both *LllS* and *RHS*. rrhen the confidence of the rule *LHS* ⇒ *RHS* is *sup(LHSU RIIS)* / *sup(LHS)*. The confidence of a rule is an indication of the strength of the rule. As an exalnple, consider again the rule *{pen}* ⇒ *{ink}*. The confidence of this rule is 75%; 75% of the transactions that contain the itenlset *{pen}* also contain the iternset *{ink}*.

26.3.2 An Algorithm for Finding Association Rules

A user can ask for all association rules that have a specified minimum support *(minsup)* and mininlum confidence *(rninconf)*, and various algorithrns have been developed for finding such rules efficiently. These algorithms proceed in two steps. In the first step, all frequent itemsets with the user-specified minimum support are computed. In the second step, rules are generated using the frequent itemsets as input. We discussed an algorithm for finding frequent iternsets in Section 26.2; we concentrate here on the rule generation part.

Once frequent iteulsets are identified, the generation of all possible candidate rules with the user-specified minirnum support is straightforward. Consider a frequent iternset X with support *sx* identified in the first step of the algorithrn. To generate a rule fronl X, we divide *X* into two iternsets, *LHS* and *RJIS*. The confidence of the rule *LllS* ⇒ *RHS* is *Sx / SLIIS,* the ratio of the support of *X* and the support of *LHS*. Frorn the a priori property, we know that the support of *LllS* is larger than *rninsup,* and thus we have C0111puted the support of *L1IS* during the first step of the algoritlnn. We can cornpute the confidence values for the candidate rule by calculating the ratio support(X)/support(LlIS) and then check how the ratio cornpares to *minconf*.

In general, the expensive step of the algorithnl is the cornputation of the frequent itenlsets, and lnany different algorithrns have been developed to perfonn this step efficiently. Rule generation –given that all frequent itcrl1sets have been identified·—·····is straightforward.

In the rest of this section, we discuss SOlne generalizations of the problern.

26.3.3 Association Rules and ISA Hierarchies

In rnany cases, an **ISA** hierarchy or category hierarchy is iInposed on the set of iterl1s. In the presence of a hierarchy, a transaction contains, for each of its iteuls, irnplicitly all the iteln's ancestors in the hierarchy. For example, consider the category hierarchy shown in Figure 26.3. Given this hierarchy, the Purchases relation is conceptually enlarged by the eight records shown in Figure 26.4. rrhat is, the Purchases relation has all tuples shown in Figure 26.1 in addition to the tuples shown in Figure 26.4.

The hierarchy allows us to detect relationships between iterns at different levels of the hierarchy. As an exarnple, the support of the itemset *{ink, juice}* is 50%, but if we replace *juice* with the more general category *beverage,* the support of the resulting itemset *{ink, beverage}* increases to 75%. In general, the support of an itemset can increase only if an item is replaced by one of its ancestors in the ISA hierarchy.

Assulning that we actually physically add the eight records shown in Figure 26.4 to the Purchases relation, we can use any algorithm for computing frequent itemsets on the augmented database. Assuming that the hierarchy fits into rnain memory, we can also perforln the addition on-the-fly while we scan the database, as an optimization.

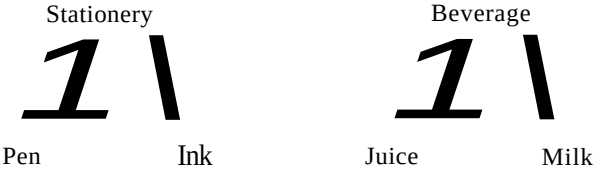


Figure 26.3 An ISA Category Taxonomy

transid	*custid*	*date*	*item*	*qty*
111	201	5/1/99	stationery	3
111	201	5/1/99	beverage	9
112	105	6/3/99	stationery	2
112	105	6/3/99	beverage	1
113	106	5/10/99	stationery	1
113	106	5/10/99	beverage	1
114	201	6/1/99	stationery	4
114	201	6/1/99	beverage	5

Figure 26.4 Conceptual Additions to the Purchases Relation with ISA Hierarchy

26.3.4 Generalized Association Rules

Although association rules have been most \videly studied in the context of market basket analysis, or analysis of cllstorner transactions, the concept is mo1'e general. Consider the Purchases relation as shown in Figure 26.5, grouped by *custid*. By exanlining the set of custorner groups, we can identify association rules such as {pen} ⇒ {rnilk}. rThis rule should now be read as follows: "If a pen is purchased by a custorner, it is likely that Inilk is also be purchased by that custcuner." In the Purchases relation shown in Figure 26.5, this rule has both support and confidence of 100%.

transid	*custid*	*date*	*item*	*qty*
112	105	6/3/99	pen	1
112	105	6/3/99	ink	1
112	105	6/3/99	milk	1
113	106	5/10/99	pen	1
113	106	5/10/99	rnilk	1
114	201	5/15/99	pen	2
114	201	5/15/99	ink	2
114	201	5/15/99	juice	4
114	201	6/1/99	water	1
111	201	5/1/99	pen	2
111	201	5/1/99	ink	1
111	201	5/1/99	rnilk	3
111	201	5/1/99	juice	6

Figure 26.5 The Purchases Helation Sorted on Customer ID

Similarly, we can group tuples by date and identify association rules that describe purchase behavior on the same day. As an exalnple consider again the Purchases relation. In this case, the rule {pen} ⇒ {rnilk} is now interpreted as follows: "On a day when a pen is purchased, it is likely that luilk is also be purchased."

If we use the *date* field as grouping attribute, we call consider a rnore general problem called calendric rnarket basket analysis. In calendric rnarket basket analysis, the user specifies a collection of calendars. A, calendar is any group of dates, such as *every Sunday in the year 1999*, or *every first of the month*. A rule holds if it holds on every day in the calendar. Given a calendar, we can cornpute a.ssociatiol1 rules over the set of tuples \vhose *date* field falls within the calendar.

By specifying interesting calendars, we can identify rules that rnight not have enough support and confidence with respect to the entire database but have enough support and confidence on the subset of tuples that fall within the calendar. On the other hand, even though a rule rnight have enough support and confidence \vith respect to the c0l11plete database, it Inight gain its support only £1'0111 tuples that fall within a calendar. In this case, the support of the rule over the tuples within the calendar is significantly higher than its support with respect to the entire database.

As an exarnple, consider the Purchases relation with the calendar *every first of the month*. \Vithin this calendar, the association rule *pen* $\Rightarrow$ *ju:ice* has support and confidence of 100%, whereas over the entire Purcha.ses relation, this rule only has 50% support. On the other hand, within the calendar, the rule *pen* $\Rightarrow$ *milk* has support of confidence of 50%, whereas over the entire Purchases relation it has support and confidence of 75%.

More general specifications of the conditions that rIlust be true within a group for a rule to hold (for that group) have also been proposed. We rnight want to say that all items in the *LHS* have to be purchased in a quantity of less than two itelTIS, and all itenls in the *RHS* rnust be purchased in a quantity of more than three.

lJsing different choices for the grouping attribute and sophisticated conditions as in the preceding exarnples, we can identify rules Inore cornplex than the basic association rules discussed earlier. These Inore cornplex rules, nonetheless, retain the essential structure of an association rule as a condition over a group of tuples, with support and confidence rneasures defined as usual.

26.3.5 Sequential Patterns

Consider the Purchases relation sho\vn in Figure 26.1. Each group of tuples, having the sarne *custid* value, can be thought of as a *sequence* of transactions ordered by *date*. rrhis allows us to identify frequently arising buying patterns over tirne.

We begin b,Y introducing the concept of a sequence of itel11sets. Each transaction is represented by a set of tuples, and by looking at the values in the *item* colurnn, we get a set of iterns purchased in that transaction. Therefore, the sequence of transactions associated with a cllstorner corresponds naturally to a sequence of itelnsets purchased by the custorner. For exalnplc, the sequence of purchases for cllstorner 201 is $\langle\{pen, ink, milk, juice\}, \{pen, ink, juice\}\rangle$.

A subsequence of a sequence of iternsets is obtained by deleting one or 11101'e itcrnsets, and is also a sequence of itenlsets. We say that a sequence $\langle a_1, \ldots, a_{rn} \rangle$ is contained in another sequence S if S has a subsequence $(b_t, \ldots, bln)$ such that $a_i \subseteq b_i$, for $1 \leq i \leq rn$. Thus, the sequence $\langle \{pen\}, \{ink, rnilk\}, \{pen, ju'ice\} \rangle$ is contained in $\langle \{pen, link\}, \{shir\cdot t\}, \{ju'ice, ink, milk\}, \{juice, pen, milk\} \rangle$. Note that the order of itenlS within each iterllset does not rnatter. However, the order of iterllsets does lllatter: the sequence $\langle \{pen\}, \{ink, rn'ilk\}, \{pen, juice\} \rangle$ is not contained in $\langle \{pen, ink\}, \{shirt\}, \{juice, pen, rnilk\}, \{juice, milk, 'ink\} \rangle$.

The support for a sequence S of iternsets is the percentage of custorner sequences of which 8 is a subsequence. The problenl of identifying sequential patterns is to find all sequences that have a user-specified rllinimurll support. A sequence $(aI, a2, a3, \ldots, am)$ with minimurn support tells us that custorners often purchase the itelns in set $a1$ in a transaction, then in sonle subsequent transaction buy the itcrlls in set $a2$, then the items in set $a3$ in a later transaction, and so on.

Like association rules, sequential patterns are staternents about groups of tuples in the current database. Cornputationally, algorithms for finding frequently occurring sequential patterns resernble algorithrns for finding frequent itemsets. Longer and longer sequences with the required rninirnum support are identified iteratively in a nlanner very similar to the iterative identification of frequent iternsets.

26.3.6 The Use of Association Rules for Prediction

Association rules are widely used for prediction, but it is inlportant to recognize that such predictive use is not justified without additional analysis or dornain knowledge. Association rules describe existing data accurately but can be misleading when used naively for prediction. For exaruple, consider the rule

$$\{pen\} \Rightarrow \{ink\}$$

The confidence associated with this rule is the conditional probability of an ink purchase given a pen purchase *over the given database;* that is, it is a *descriptive* rueasure. We rnight use this rule to guide future sales prornotions. For exalllple, we rnight offer a discount on pens to increase the sales of pens and, therefore, aIso increase sales of ink.

Flowever, such a prorllotion assumes that pen purchases are good indicators of ink purchases in *future* custC)Iuer transactions (in addition to transactions in the current database). This assumption is justified if there is a *causal link* between pen purchases and ink purchases; that is, if buying pens causes the buyer to also buy ink. Ifowever,we can infer association rules\vith high support

and confidence in SOlnc situations where there is no causal link between *L118* and *RIIS*. For exarnple, suppose that pens are ahvays purchased together with pencils, perhaps because of customers' tendency to order writing instrulllents together. We would then infer the rule

$$\{pencil\} \Rightarrow \{ink\}$$

with the saBle support and confidence as the rule

$$\{pen\} \Rightarrow \{ink\}$$

However, there is no causal link between pencils and ink. If we prornote pencils, a custolner who purchases several pencils due to the pronlotion has no reason to buy Inore ink. Therefore, a sales prolnotion that discounted pencils in order to increase the sales of ink would fail.

In practice, one would expect that, by exallllnlng a large database of past transactions (collected over a long tirne and a variety of circumstances) and restricting attention to rules that occur often (i.e., that have high support), we rninirnize inferring lnisleading rules. However, we should bear in rnind that nlisleading, noncausal rules lnight still be generated. Therefore, we should treat the generated rules as possibly, rather than conclusively, identifying causal relationships. Although association rules do not indicate causal relationships between the *LHS* and *RHS,* we elllphasize that they provide a useful starting point for identifying such relationships, using either further analysis or a dornain expert's judgrnent; this is the reason for their popularity.

26.3.7 Bayesian **Networks**

Finding causal relationships is a challenging task, as we saw in Section 2G.3.6. In general, if certain events are highly correlated, there are rnany possible explanations. For exalnple, suppose that pens, pencils, and ink are purchased together frequently. It rnight be that the purchase of one of these itelIlS (e.g., ink) depends causally on the purchase of another itern (e.g., pen). ()r it Blight be that the purchase of one of these iterns (e.g., pen) is strongly correlated with the purchase of another (e.g., pencil) because of sorne underlying phenornenon (e.g., users' tendency to think about \vriting instrulnents together) that causally influences both purchases. How can we identify the true causal relationships that hold between these events in the real world?

One approach is to consider each possible cOlnbination of causal relationships arnong the varial)les or events of interest to us and evaluate the likelihood of each cornbination on the basis of the data available to us. If we think of each cornbination of causal relationships as a *model* of the real world underlying the

collected data, we can assign a score to each ruode! by considering how consistent it is (in terms of probabilities, 'with senne sin1plifying assumptions) with the observed data. Bayesian networks are graphs that can be used to describe a class of such Il1odels, with one node per variable or event, and arcs between nodes to indicate causality. For exarnpIe, a good Iuodel for our running exarnpIe of pens, pencils, and ink is shown in Figure 26.6. In general, the nurnber of possible Inodels is exponential in the nurnber of variables, and considering all rnodels is expensive, so SOUle subset of all possible rnodels is evaluated.

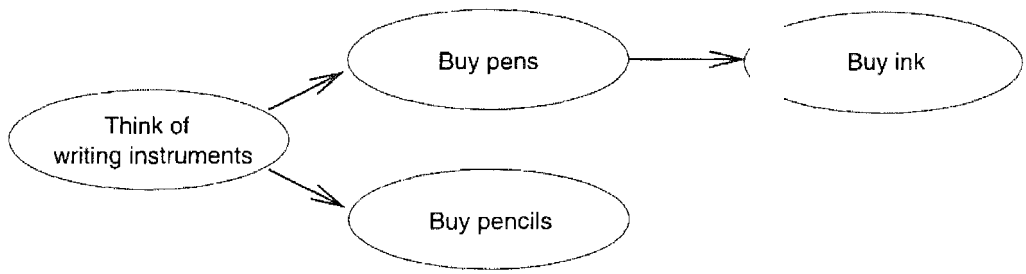


Figure 26.6 Bayesian Network Showing Causality

26.3.8 Classification and Regression Rules

Consider the following view that contains inforrnation froln a rnailing carnpaign perforrned by an insurance cornpany:

InsuranceInfo(*age:* integer, *cartype:* string, *highrisk:* boolean)

The Insurancelnfo vie\v has inforrnation about current cllston1ers. Each record contains a cllstolner's age and type of car as well as a flag indicating whether the person is considered a high-risk custorner. If the flag is true, the cllstorner is considered high-risk. We would like to use this information to identify rules that predict the insurance risk of new insurance applicants whose age and car type are known. :For exarnple, one such rule could be: "If *age* is between 16 and 25 and *cartype* is either Sports or.Truck, then the risk is high."

Note that the rules we want to find have a specific structure.vVe are not interested in rules that predict the age or type of car of a person: we are interested only in rules that predict the insurance risk. Thus, there is one designated attribute whose value we wish to predict, and we call this attribute the dependent attribute. rrhe other attributes aTe called predictor attributes. In our example, the dependent attribute in the Insurancelnfo vic\v is the *highrisk* attribute arld the predictor attributes are *age* and *cartype*. The general foru1 of the types of rules we want to discover is

$$P_1(X_1) \wedge P_2(X_2) \ldots \wedge P_k(X_k) \Rightarrow Y = c$$

The predictor attributes $X_1, \ldots, X_k$ are used to predict the value of the dependent attribute Y. Both sides of a rule can be interpreted as conditions on fields of a tuple. The $Pi(X_i)$ are predicates that involve attribute X_i. The fornl of the predicate depends on the type of the predictor attribute. We distinguish two types of attributes: numerical and categoricaL For numerical attributes, we can perfOrIn nurnerieal cornputations, such as cornputing the average of two values; whereas for categorical attributes, the only allowed operation is testing whether two values are equal. In the InsuranceInfo view, *age* is a nUlllerical attribute whereas *cartype* and *highrisk* are categorical attributes. Returning to the forrn of the predicates, if X_i is a nUlllerical attribute, its predicate P_i is of the forln $li \leq X_i \leq hi$; if X_i is a categorical attribute, *Pi* is of the forIll $X_i \in \{Vl, \ldots, Vj\}$.

If the dependent attribute is categorical, we call such rules classification rules. If the dependent attribute is nurnerical, we call such rules regression rules.

For exarnple, consider again our exaInple rule: "If *age* is between 16 and 25 and *caTtype* is either Sports or Truck, then *highrisk* is true." Since *highrisk* is a categorical attribute, this rule is a classification rule. We can express this rule fonnally as follows:

$$(16 \leq age \leq 25) \wedge (cartype \in \{\text{Sports, Truck}\}) \Rightarrow highri8k = \text{true}$$

We can define support and confidence for classification and regression rules, as for association rules:

- Support: ffhe support for a condition C is the percentage of tuples that satisfy C. The support for a rule $C1 \Rightarrow C2$ is the support for the condition $C1 \wedge C2$.
- Confidence: Consider those tuples that satisfy condition $C1$. The confidence for a rule $C1 \Rightarrow C2$ is the percentage of such tuples that also satisfy condition $C2$.

As a further generalization, consider 1,118 right-hand side of a classification or regression rule: Y =. c..Each rule predicts a value of Y for a given tuple based on the values of predictor attributes X1, ... ,Xk. We can consider rules of the fonn

$$P_1(X_1) \wedge \ldots \wedge P_k(X_k) \Rightarrow Y = f(X_1, \ldots, X_k)$$

where f is sonlC function. We do not discuss such rules further.

Classification and regression rules differ fr0111 clssociation rules by considering continuous and categorical fields, rather than only one field that is set-valued. Identifying such rules efficiently presents a new set of challenges; we do not

discuss the general case of discovering such rules. We discuss a special type of such rules in Section 26.4.

Classification and regression rules have many applications. Exarnples include classification of results of scientific experirnents, where the type of object to be recognized depends on the measurements taken; direct lllail prospecting, where the response of a given customer to a prolnotion is a function of his 01' her inCOlue level and age; and car insurance risk assesslnent, where a customer could be classified as risky depending on age, profession, and car type. Example applications of regression rules include financial forecasting, where the price of coffee futures could be SOlne function of the rainfall in Colornbia a month ago, and Inedical prognosis, where the likelihood of a tUInor being cancerous is a function of Illeasured attributes of the tUlnor.

26.4 TREE·STRUCTURED RULES

In this section, we discuss the problem of discovering classification and regression rules from a relation, but we consider only rules that have a very special structure. The type of rules we discuss can be represented by a tree, and typically the tree itself is the output of the data mining activity. Trees that represent classification rules are called classification trees or decision trees and trees that represent regression rules are called regression trees

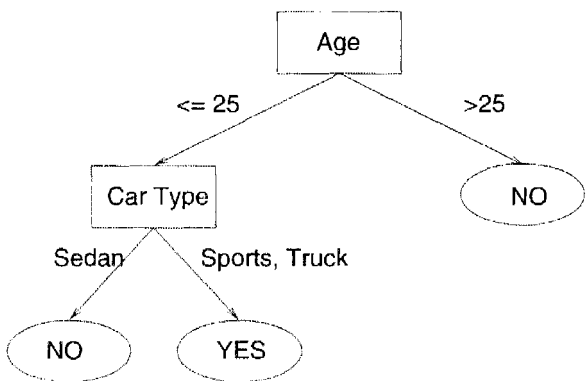


Figure 26.7 Insurance Risk Example Decision Tree

A.s an exalnple, consider the decision tree ShO\Vll in Figure 26.7. Each path froln the root node to a leaf node represents one classification rule. For example, the path fron1 the root to t11e leftrnost leaf node represents the classification rule: "If a person is 25 years or .younger and drives a sedan, then he or she is likely to have a low insurance risk." The path fforn the root to the right-Inost leaf node represents the classification rule: "If a person is older than 25 years, then he or she is likely to have a low insurance risk."

Tree-structured rules are very popular since they are easy to interpret. Ease of understanding is very hllportant because the result of any data rninillg activity needs to be cOlllprehensible by nonspecialists. In addition, studies have shown that, despite Ihnitations in structure, tree-structured rules are very accurate. There exist efficient algorithrl1s to construct tree-structured rules fronl large databases. We discuss a sample algorithrIl for decision tree construction in the rernainder of this section.

26.4.1 Decision Trees

A decision tree is a graphical representation of a collection of classification rules. Given a data record, the tree directs the record frOIn the root to a leaf. Each internal node of the tree is labeled with a predictor attribute. This attribute is often called a splitting attribute, because the data is 'split' based on conditions over this attribute. The outgoing edges of an internal node are labeled with predicates that involve the splitting attribute of the node; every data record entering the node must satisfy the predicate labeling exactly one outgoing edge. T'he cornbined information about the splitting attribute and the predicates on the outgoing edges is called the splitting criterion of the node. A node with no outgoing edges is called a leaf node. Each leaf node of the tree is labeled with a value of the dependent attribute. We consider only binary trees where internal nodes have two outgoing edges, although trees of higher degree are possible.

Consider the decision tree shown in Figure 26.7. The splitting attribute of the root node is *age,* the splitting attribute of the left child of the root node is *cartype.* The predicate on the left outgoing edge of the root node is $age \leq 25$, the predicate on the right outgoing edge is $age > 25$.

"'e can no\v aBsociate a classification rule with each leaf node in the tree as follows. Consider the path frorH the root of the tree to the leaf node. Each edge on that path is labeled with a predicate. 'The conjunction of all these predicates rnakes up the left-hand side of the rule. rrhe value of the dependent attribute at the leaf node rnakesup the right-ha,nd side of the rule. Thus, the deeision tree represents a collection of classification rules, OIle for each leaf node.

A decision tree is usually constructed in t\VO phases. In phase one, the growth phase, an overly large tree is constructed. This tree represents the records in the input database very accurately; for exaluple, the tree rnight contain leaf nodes for inclividual records frorn the input database. Tn phase t\VO, the pruning phase, the final size of the tree is deterrnined. The rules represented by the tree constructed in phase one are usuall:y overspecialized. By reducing the size of the tree, we generate a srnaller nUlnber of lllore general rules that

are better than a very large nUlllbcr of very specialized rules. Algorithrns for tree pruning are beyond our scope of discussion here.

Classification tree algorithrlls build the tree greedily top-down in the following way. At the root node, the database is exarnined and the locally 'best' splitting criterion is cornputed. rrhe database is then partitioned, according to the root node's splitting criterion, into two parts, one partition for the left child and one pa,rtition for the right child. The algoritlull then recurses on each child. rrhis schcrua is depicted in Figure 26.8.

Input: !loden, partition *D*, split selection ruethod *S*
Output: decision tree for D rooted at node *n*

Top-Down Decision Tree Induction Schema:
BuildTree(Node 11, data partition *D*, split selection rnethod *S)*
(1) Apply *S* to *D* to find the splitting criterion
(2) if (a good splitting criterioll is found)
(3) Create two children nodes n_1 and *n2* of *n*
(4) Partition *D* into D_1 and *D2*
(5) BuildT'ree(nl, D_1, *S)*
(6) Build'Tree(n2, *D2*, *S)*
(7) endif

Figure 26.8 Decision Tree Induction Schema

The splitting criterion at a node is found through application of a split selection method. A split selection rnethod is an algorithrIl that takes as input (part of) a relation and outputs the locally 'best' splitting criterion. In our exarnple, the split selection rnethod exarnines the attributes *cartype* and *age,* selects one of thern as splitting attribute, and then selects the splitting predicates. IVlany different, very sophisticated split selection rnethods have been developed; the references provide pointers to the relevant literature.

26.4.2 An Algorithm to Build Decision Trees

If the input database fits into rna,in Inernory, we can directly follow th.e classification tree induction schcrna shown in Figure 26.8. How can we construct decision trees when the input relation is larger than rnain rncrJlory? In this case, step (1) in Figllre 26.8 fails, since the input database does not fit in Inenl0ry. But we can rnake one irnportant observation about split selection Inethods that helps us to reduce the rnain rnerllory requircluents.

Consider a node of the decision tree. The split selection rnethod has to Inake two decisions after exarllining the partition at that node: It has to select the splitting attribute, and it has to select the splitting predicates for tIle outgo-

age	*cartype*	*highrisk*
23	Sedan	false
30	Sj)orts	false
36	Sedan	false
25	Truck	true
30	Sedan	false
23	Truck	true
30	Truck	false
25	Sports	true
18	Sedan	false

Figure 26.9 The Insurancelnfo Relation

ing edges. After selecting the splitting criterion at a node, the algorithrn is recursively applied to each of the children of the node. Does a split selection rnethod actually need the cornplete database partition as input? Fortunately, the answer is no.

Split selection rnethods that cornpute splitting criteria that involve a single predictor attribute at each node evaluate each predictor attribute individually. Since each attribute is exarnined separately, we can provide the split selection rnethod with aggregated inforulation about the database instead of loading the cornplete database into rnain rnenlory. Chosen correctly, this aggregated inforrnation enables us to cornpute the same splitting criterion as we would obtain by exarnining the conlplete database.

Since the split selection rnethod exanlines all predictor attributes, we need aggregated inforrnation about each predictor attribute. We call this aggregated inforrnation the AVe set of the predictor attribute. The AVe set of a predictor attribute X at noden is the projection of n's database partition onto X and the dependent attribute where counts of the individual values in the dorllain of the dependent attribute are aggregated. (AVC stands for Attribute-Value, Class label, because the values of the dependent attribute are ofterl called class labels.) For example, consider the Insurancelnfo relation as shown in Figure 26.9. rrhe AVe set of the root node of the tree for predictor attribute *age* is the result of the following database query:

```
SELECT    R.age, Il.highrisk, COUNT (*)
FROM      Insurancelnfo R
GROUP BY  R.age, R.highrisk
```

The AVe set for the left child of the root node for predictor attribute *cartype* is the result of the following query:

```
SELECT    R.cartype, R.highrisk, COUNT (*)
FROM      InsuranceInfo R
WHERE     R.age <= 25
GROUP BY  R.cartype, R.highrisk
```

The two AVC sets of the root node of the tree are shown in Figure 26.10.

Car type	highrisk	
	true	false
Sedan	0	4
Sports	1	1
Truck	2	1

Age	highrisk	
	true	false
18	0	1
23	1	1
25	2	0
30	0	3
36	0	1

Figure 26.10 AVC Group of the Root Node for the InsuranceInfo Relation

We define the AVC group of a node n to be the set of the AVC sets of all predictor attributes at node n. Our example of the InsuranceInfo relation has two predictor attributes; therefore, the AVC group of any node consists of two AVC sets.

How large are AVC sets? Note that the size of the AVC set of a predictor attribute X at node n depends only on the number of distinct attribute values of X and the size of the domain of the dependent attribute. For example, consider the AVC sets shown in Figure 26.10. The AVC set for the predictor attribute *cartype* has three entries, and the AVC set for predictor attribute *age* has five entries, although the InsuranceInfo relation as shown in Figure 26.9 has nine records. For large databases, the size of the AVC sets is independent of the number of tuples in the database, except if there are attributes with very large domains, for example, a real-valued field recorded at a very high precision with many digits after the decimal point.

If we make the simplifying assumption that all the AVC sets of the root node together fit into main memory, then we can construct decision trees for very large databases as follows: We make a scan over the database and construct the AVC group of the root node in memory. Then we run the split selection method of our choice with the AVC group as input. After the split selection method computes the splitting attribute and the splitting predicates on the outgoing nodes, we partition the database and recurse. Note that this algorithm is very similar to the original algorithm shown in Figure 26.8; the only modification necessary is shown in Figure 26.11. In addition, this algorithm is still independent of the actual split selection method involved.

Input: node n, partition D, split selection 111ethod S
Output: decision tree for D rooted at node n

Top-Down Decision Tree Induction Schenla:
BuHdTree(Node n, data partition D, split selection method S)
(1a) Make a scan over D and construct the AVe group of n in-nlCIllory
(1b) Apply S to the AVe group to find the splitting criterion

Figure 26.11 Classification m-ee Induction Refinement with AVe Groups

26.5 CLUSTERING

In this section we discuss the clustering problem. The goal is to partition a set of records into groups such that records within a group are shnilar to each other and records that belong to two different groups are dissimilar. Each such group is called a cluster and each record belongs to exactly one cluster.[1] Sirnilarity between records is Ineasured cOlnputationally by a distance function. A distance function takes two input records and returns a value that is a measure of their silnilarity. Different applications have different notions of similarity, and no one rneasure works for all domains.

As an exarnple, consider the scherna of the Custol11erlnfo view:

CustornerInfo(*age:* int, *salary:* real)

We can plot the records in the view on a two-dil11ensional plane as shown in Figure 26.12. The two coordinates of a record are the values of the record's *salary* and *age* fields. \Ve can visually identify three clusters: Young cllstorners who have low salaries, young cllstorners with high salaries, and older cnstorners with high salaries.

Usnally, the output of a clustering algorithrll consists of a, summarized representation of each cluster. The type of sUIrllnarized representation depends strongly on the type and shape of clusters the algoritlull cornputes. For exarnple, assume that we have spherical clusters as in the exalllple shown in Figure 26.12. We can surnrnarize each cluster by its *center* (often also called the *mean*) and its *radius*, which are defined as follo\vs. Given a collection of records $r_1, \ldots, r_n$, their center C and radius R are defined as follows:

$$C = \frac{1}{n}\sum_{i=1}^{n} r_i, \quad R = \sqrt{\frac{\sum_{i=1}^{n}(r_i - C)}{n}}$$

[1] There are clustering algorithrns that allow overlapping clusters, where a record could belong to several clusters.

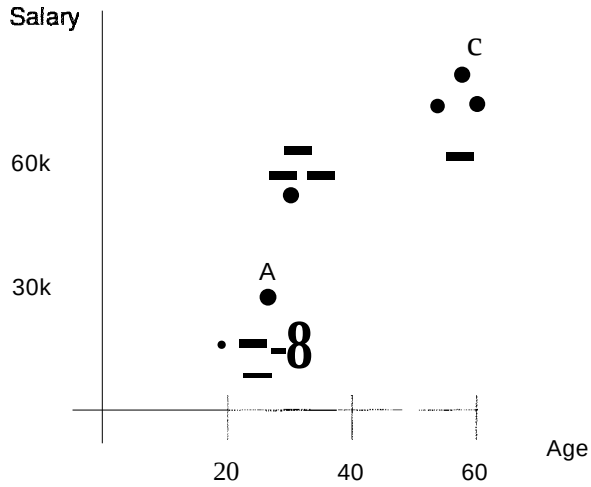


Figure 26.12 Records in CustomerInfo

There are two types of clustering algorithlns. A **partitional** clustering algorithnl partitions the data into *k* groups such that SOUle criterion that evaluates the clustering quality is optirnized. The nurnber of clusters *k* is a parameter whose value is specified by the user. A **hierarchical** clustering algorithnl generates a sequence of partitions of the records. Starting with a partition in which each cluster consists of one single record, the algorithrn rnerges two partitions in each step until only one single partition rernains in the end.

26.5.1 A Clustering Algorithm

Clustering is a very old problern, and nurnerous algorithnls have been developed to cluster a collection of records. Traditionally, the nurnber of reeords in the input database was assurned to be relatively slnall and the cornplete database was assurned to fit into Inain rnernory. In this section,we describe a clustering algoritlnn called BIRCH that handles very large databases. The design of BIR,CII reflects the followving two assumptions:

- The rnunber of records is potentially very large, and therefore we want to rnake only one scan over the database.
- Only a lirnited arnount of rnain rnenlory is available.

A user can set t\VO pararneters to control the BIRCH algoritllln. The first is a thresl10lcl on the arnount of rnain luernory available. This main rncrnory threshold translates into a lllaxirnurn nurnber of cluster SUIJlrnaries *k* that can be maintained in rncrllory. 'The second pararneter ϵ is an initial threshold for the radius of an,Y cluster. The value of ϵ is an upper bound on the radius of any cluster and controls the nUlnber of clusters that the algorithrn discovers. If ϵ is slnalI, we discover many sInaII clusters; if ϵ is large, we discover very fe\v

clusters, each of which is relatively large. We say that a cluster is compact if its radius is s1na11e1' than ϵ.

BIRCH always lTlaintains k or fewer cluster sU1nrnaries (C_i, R_i) in rnain Hlcnl0ry, where C_i is the center of cluster i and R_i is the radius of cluster i. The a1gorithrn ahvays rnaintains cornpact clusters; that is, the radius of each cluster is less than E. If this invariant cannot be rnaintained with the given arIlount of rnain lnerno1'y, ϵ is increased as described next.

The algoritlnl1 reads records frorn the database sequentially and processes the1l1 as follows:

1. Cornpute the distance betVileen record r and each of the existing cluster centers. Let i be the cluster index such that the distance between rand C_i is the srnallest.
2. Cornpute the value of the new radius R_i' of the ith cluster under the assumption that r is inserted into it. If $R_i' \leq \epsilon$, then the ith cluster rernains cornpact, and we assign r to the ith cluster by updating its center and setting its radius to R_i'. If $R_i' > \epsilon$, then the ith cluster would no longer be cOlnpact if we insert r into it. Therefore, we start a new cluster containing only the record r.

The second step presents a problern if we already have the rnaxinnun nurnber of cluster sUIInnaries, k. If we now read a record that requires us to create a new cluster, we lack the rnain rne1nory required to hold its surnrnary. In this case, we increase the radius threshold E-----using SOHle heuristic to detennine the increase---in order to *merge* existing clusters: An increase of ϵ has two consequences. First, existing clusters can accorl1rnodate rnore records, since their rnaxirnurn radius has increased. Second, it Blight be possible to rnerge existing clusters such that the resulting cluster is still cornpact. rrhus, an increase in ϵ usually reduces the l1ulIlber of existing clusters.

The cornplete BIRCH algorithrl1 uses a balanced in-rnernory tree, which is sirnilar to a B+ tree in structure, to quickly identify the closest cluster center for a new record. A description of this data structure is beyond the scope of our discussion.

26.6 SIMILARITY SEARCH OVER SEQUENCES

A lot of inforrnation stored in datal)ases consists of sequences. In this section, we introduce the problern of siInilarity search over a collection of sequences. Our query Inode} is very sirnple: We assurne that the user specifies a query sequence and wants to retrieve all data sequences that are similar to the

Commercial Data Mining Systems: There area number of data ruining products on the rnarket today, such as SASEnterprise Miner, SPSS Clenlcntine, CART froIn Salford Systems, Megaputer PolyAnalyst, ANGOSS I<nowledgeStudio. We highlight two that have strong database ties.

IBM's Intelligent Miner offers a wide range of algorithlns, including association rules, regression, classification, and clustering. The emphasis of Intelligent Miner is on scalability—the product contains versions of all algorithllls for parallel cOlnputers and is tightly integrated with IBM's DB2 database systenl. DB2's object-relational capabilities can be used to define the data Inining classes of SQL/MM. Of course, other data 111ining vendors can use these capabilities to add their own data mining models and algorithms to DB2.

Microsoft's SQL Server 2000 has a component called the Analysis Server that lnakes it possible to create, apply, and lnanage data mining models within the DBMS. (SQL Server's OLAP capabilities are also packaged in the Analysis Server component.) The basic approach taken is to represent a mining rrlodel as a table; clustering and decision tree models are currently supported. The table conceptually has one row for each possible combination of input (predictor) attribute values. The model is created using a staternent analogous to SQL's CREATE TABLE that describes the input on which the model is to be trained and the algorithrn to use in constructing the model. An interesting feature is that the input table can be defined, using a specialized view rnechanisnl, to be a *nested table.* For exalnple,we can define an input table with one row per custolner, where one of the fields is a nested table that describes the eustolner's purchases. The SQL/MM extensions for data ruining do not provide this capability because SQL:1999 does not currently support nested tables (Section 23.2.1). Several properties of attributes, such as whether they are discrete or continuous, can also be specified.

A rnodel is trained by inserting rows into it, using the INSERT cornlnand. It is applied to a new dataset to lnake predictions using a new kind of join called PREDICTION JOIN; in principle, each input tuple is matched with the corresponding tuple in the rnining lllodel to detennine the value of the predicted attribute. Thus, end users can create, train,and apply decision trees and clustering using extended SQL. 'There are also cornrnands to browse rnodels. Unfortnnately, users cannot add new rnodels or new algorithrlns for models, a capability that is supported in the SQL/MlVI proposa.

query sequence. Sinlilarity search is different frorH 'normal' queries in that we are interested not only in sequences that rnatch the query sequence exactly but also those that differ only slightly frorn the query sequence.

We begin by describing sequences and sirnilarity between sequences. A data sequence X is a series of nurnbers X = $\langle x_1, \ldots Xk \rangle$. S0111etirIles X is also called a time series. We call k the length of the sequence. A subsequence Z = (Zl' ... , $z_j\rangle$ is obtained frolll another sequence X = $(Xl, \ldots, Xk)$ by deleting nurnbers froln the front and back of the sequence X. ForInally, Z is a subsequence of X if $z_1 = x_i$, $z_2 = x_{i+1}$, $z_j = z_{i-tj-1}$ for SaIne i E {I, ... , $k-j+$I}. Given two sequences X = (Xl, Xk) and Y = $(Yll'' \ldots, Yk)$, we can define the Euclidean **Darrri** as the distance between the two sequences as follows:

$$\|X - Y\| = \sum_{i=1}^{k} (x_i - y_i)^2$$

Given a user-specified query sequence and a threshold pararneter ϵ, our goal is to retrieve all data sequences that are within E-distance of the query sequence.

Sirnilarity queries over sequences can be classified into two types.

- Complete Sequence Matching: The query sequence and the sequences in the database have the sarne length. Given a user-specified threshold paranleter ϵ, our goal is to retrieve all sequences in the database that are within E-distance to the query sequence.
- Subsequence Matching: rrhe query sequence is shorter than the sequences in the database. In this case, we want to find all subsequences of sequences in the database such that the subsequence is within distance ϵ of the query sequence. We do not discuss subsequence rnatching.

26.6.1 An Algorithm to Find Similar Sequences

Given a collection of data sequences, a query sequence, and a distance threshold ϵ, how can we efficiently find all sequences within f-distance of the query sequence?

()ne possibility is to scan the database, retrieve each data sequence, and C0111pute its distance to the query sequence. \While this algorithrn has the rnerit of being sirnple, it always retrieves every data sequence.

Because we consider the conlplete sequence lnatehing problenl, all data sequences and the query sequence have the same length. We can think of this sirnilarity search as a high-dirnensional indexing probleul. Each data sequence

and the query sequence can be represented as a point in a k-dirnensionaJ space. Therefore, if we insert all data sequences into a Illuitidirnensional index, we can retrieve data sequences that exactly ll1atch the query sequence by qllerying the index. But since 'we want to retrieve not only data sequences that Inatch the query exactly but also all sequences within (-distance of the query sequence, we do not use a point query as defined by the query sequence. Instead, we query the index 'with a hyper-rectangle that has side-length 2E and the query sequence as center, and we retrieve all sequences that fall within this hyper-rectangle. We then discard sequences that are actually further than ϵ away froln the query sequence.

ITsing the index allows us to greatly reduce the nurnber of sequences we consider and decreases the time to evaluate the sirnilarity query significantly. The bibliographic notes at the end of the chapter provide pointers to further improvernents.

26.7 INCREMENTAL MINING· AND DATA STREAMS

Real-life data is not static, but is constantly evolving through additions or deletions of records. In sorne applications, such as network Inonitoring, data arrives in such high-speed strearns that it is infeasible to store the data for offline analysis. We describe both evolving and strearning data in terlns of a framework called block evolution. In block evolution, the input dataset to the data mining process is not static but periodically updated with a new block of tuples, for exarnple, every day at rnidnight or in a continuous strealn. A block is a set of tuples added siInultaneously to the database. For large blocks, this Inodel captures comrnon practice in rnany of today's data warehouse installations, where updates from operational databases are batched together and perforrned in a block update. For srnall blocks of data—at the extrerne, each block consists of a single record—this rnodel captures strealning data.

In the block evolution rnodel, the database consists of a (conceptually infinite) sequence of data blocks $D_1, D_2, \ldots$ that arrive at tilnes I, 2, ..., \Vhe1'8 each block D_i consists of a set of records.[2] We call i the *block identifier* of block B_i. Therefore, at any titHe t, the database consists of a finite sequence of blocks of data $\langle D_1, \ldots, D_t \rangle$ that arrived at tirnes $\{1, 2, \ldots, t\}$. The database at tilne t, \vhic.h we denote by $D[1, t]$, is the union of the database at time t - 1 and the block that arrives at tirue t, D_t.

For evolving data, two classes of problerns are of particular interest: rnodel Inaintenance and change detection. The goal of 1110del maintenance is to

2In general, a block specifies records to change or delete, in addition to records to insert. We only consider inserts.

maintain a data rnining ulodcl under insertion and deletions of blocks of data. To incrernentally cornpute the data mining rnodel at time t, which we denote by $M(D[1,t))$, we HUlst consider only $M(D[1,t-1])$ and $.D_t$; we cannot consider the data that arrived prior to time t. Further, a data analyst rllight specify tirne-dependent subsets of $D[1,t]$, such as a window of interest (e.g., all the data seen thus far or last week's data). More general selections are also possible, for exarnple, all weekend data over the past year. Given such selections, we Hlllst incrernentally CCHupute the rnodel on the appropriate subset of $.D[l,t]$ by considering only $[J_t$ and the model on the appropriate subset of $1)[1,t-1]$. 'Alrnost' incrernental algoritlulls that occasionally exarnine older data rnight be acceptable in warehouse applications, where incrementality is lTIotivated by efficiency considerations and older data is available to us if necessary. This option is not available for high-speed data strearns, where older data may not be available at all.

The goal of change detection is to quantify the difference, in terrns of their data characteristics, between two sets of data and determine whether the change is rneaningful (i.e., statistically significant). In particular, we rnust quantify the difference between the rllodels of the data as it existed at sonle time tI and the evolved version at a subsequent ·tirne $t2$; that is, we Blust quantify the difference between $1\backslash 1(D[I,tI])$ and $1\backslash 1(D[1,t2])$. We can also measure changes with respect to selected subsets of data. Several natural variants of the problem exist; for exarnple, the difference between $M(D[I,t-1])$ and $M(D_t)$ indicates whether the latest block differs substantially frorn previously existing data. In the rest of this chapter, we focus on rnodel rnaintenance and do not discuss change detection.

Incrernental rnodel rnaintenance has received rnuch attention. Since the quality of the data rllining rnodel is of utrnost irnportance, incrernental rnodel rnaintena,nce a1gorithrns have concentrated on cornputing exactly the sarne Inode1 as cOlnputed by running the basic rnodel construction algoritlul1 on the union of old and new data. ()ne \videly used scalability technique is localization of changes due to new blocks. For exarnple, for density-based clustering algoritluns, the insertion of a new record affects only clusters in the neighborhood of the record, and thus efficient algorithrlls can *localize* the change to a few clusters and avoid reccHnputing all clusters. As another exarllple, in decision tree construction, we rnight be able to show that the split criterion at a, node of the tree changes only within acceptably srnall confidence intervals when records are inserted, if we assume tha,t the underlying distribution of training records is static.

One-pass rnodel construction over data strearllS has received particular attention, since data arrives and rnust be processed continuously in several ernerging application dCHnains. For exarnple, network installations of large TelecOll1

and Internet service providers have detailed usage inforruation (e.g., eall-detail-records, router packet-flow and trace data) froln different parts of the underlying network that needs to be continuously analyzed to detect interesting trends. Other exanlples include webserver logs, streaI11S of transactional data froI11 large retail chains, and financial stock tickers.

When working with high-speed data strearlls, algoritlulls lllUSt be designed to construct data rnining rnodels while looking at the relevant data iterrlS *only once and in a .fixed order* (deternlined by the strearn-arrival pattern), with a lirnited arnount of main 111eI1l0ry. Data-strearn coruputatioll has given rise to several recent (theoretical and practical) studies of online or one-pass algorithrlls with bounded HIeIIIory. Algorithrns have been developed for one-pass cornputation of quantiles and order-statistics, estirnation of frequency I1l0Inents and join sizes, clustering and decision tree construction, estimating correlated aggregates, and cOInputing one-dirnensional (i.e., single-attribute) histogranls and 1Iaa1' wavelet decolllpositions. Next, we discuss one such algorithIn, for incremental rnaintenance of frequent itemsets.

26.7.1 Incremental Maintenance of Frequent Itemsets

Consider the Purchases Relation shown in Figure 26.1 and assurne that the minimum support threshold is 60%. It can be easily seen that the set of frequent iternsets of size 1 consists of $\{pen\}$, *{ink},* and *{rnilk}* with supports of 100%, 75%, and 75%, respectively. T'he set of frequent itell1Sets of size 2 consists of $\{pen, ink\}$ and $\{pen, milk\}$, both with supports of 75%. The Purchases relation is our first block of data. Our goal is to develop an algorithrIl that rnaintains the set of frequent itcrl1sets under insertion of new blocks of data.

As a first exarnple, let us consider the addition of the block of data shown in Figllre 26.13 to our original database (Figure 26.1). V'nder this addition, the set of frequent itcrIlsets does not change, although their support values do: *{pen}, {i'nk},* and $\{milk\}$ now have support values of 100%, 60%, and 60%, respectively, and *{pen, ink}* and $\{pen, milk\}$ now have 60% support. Note that we could detect this case of 'no change' sirnply by rnaintaining the nurnber of rnarket baskets in which each iternset occured. Irl this example, we update the (al)solute) support of itcrnset *{pen}* by 1.

transid	*custid*	*date*	*item*	*qty*
115	201	7/1/99	pen	2

Figure 26.13 The Purchases Relation Block 2

transid	*custid*	*date*	*item*	*qty*
115	201	7/1/99	water	1
115	201	7/1/99	lllilk	1

Figure 26.14 The Purchases Relation Block 2a

In general, the set of frequent itemsets Illay change. As an exalnple, consider the addition of the block shown in Figure 26.14 to the original database shown in Figure 26.1. We see a transaction containing the itern water, but we do not know the support of the iterllset $\{water\}$, since water was not above the InininUlm support in our original database. A sirnple solution in this case is to rnake an additional scan over the original database and cornpute the support of the itenlset *{water}.* But can we do better? Another innnediate solution is to keep counters for *all* possible iterllsets, but the nUlnber of all possible itemsets is exponential in the nurnber of iterns-·-and most of these counters would be 0 anyway. Can we design an intelligent strategy that tells us *which* counters to ruaintain?

We introduce the notion of the negative border of a set of iternsets to help decide which counters to keep. The negative border of a set of frequent itemsets consists of all iterllsets X such that X itself is not frequent, but all subsets of X are frequent. For example, in the case of the database shown in Figure 26.1, the following iternsets rnake up the negative border: *{juice},* *{water},* and *{ink, milk}.* Now we can design a more efficient algorithm for maintaining frequent iternsets by keeping counters for all currently frequent iternsets *and* all iterllsets currently in the negative border. ()nly if an iternset in the negative border becomes frequent do we need to read the original dataset again, to find the support for new candidate itemsets that Blight be frequent.

We illustrate this point through the following tvo exarnples. If we add Block 2a shown in Figure 26.14 to the original database shown in Figure 26.1, we increase the support of the frequent iterllset $\{milk\}$ by one, and we increase the support of the iternset *{water},* which is in the negative border, by one as well. But since no iternset in the negative border beearne frequent, we do not have to re-scan the original database.

In eontrast, consider the addition of Block 2b shown in Figure 26.15 to the original database shown in Figure 26.1. In this case, the iternset *{juice},* which was originally in the negative border, becornes frequent with a support of 60%. rrhis rneans that now the following itcrnsets of size two enter the negative border: $\{juice, pen\}$, *{juice, ink},* and $\{juice, milk\}$. (We know that $\{juice, water\}$ cannot be frequent since the iteulset $\{water\}$ is not freqlient.)

115	201	7/1/99	juice	2
115	201	7/1/99	water	2

Figure 26.15 The Purchases Relation Block 2b

26.8 ADDITIONAL DATA MINING rfASKS

We focused on the problern of discovering patterns frorn a database, but there are several other equally inlportant data ruining tasks. We now discuss SOIIIe of these briefly. The bibliographic references at the end of the chapter provide luauy pointers for further study.

- Dataset and Feature Selection: It is often irnportant to select the 'right' dataset to rnine. Dataset selection is the process of finding which datasets to uline. Feature selection is the proeess of deciding which attributes to include in the mining process.
- Sampling: One way to explore a large dataset is to obtain one or luore *samples* and analyze them. The advantage of sampling is that we can carry out detailed analysis on a sarnple that would be infeasible on the entire dataset, for very large datasets. The disadvantage of sampling is that obtaining a representative salllple for a given task is difficult; we rnight rniss irnportant trends or patterns because they are not reflected in the san1ple. Current database systerns also provide poor support for efficiently obtaining sanlples. Irnproving database support for obtaining sarnples with various desirable statistical properties is relatively straightforward and likely to be available in future DBMSs. Applying sarnpling for data ruining is an area for further research.
- Visualization: Visualization techniques can significantly assist in understanding cornplex datasets and detecting interesting patterns, and the importance of visualization in data ruining is widely recognized.

26.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What is the role of data rnining in theKI)I) process? (Secti.on 26.1)
- What is the a priori property? I)escribe an algorithnl for firlding frequent itcrIlsets. (Section 26.2.1)

- How are iceberg queries related to frequent iterIlsets? (Section 26.2.2)
- Give the definition of an *association rule*. What is the difference between support and confidence of a rule? (Setion 26.3.1)
- Can you explain extensions of association rules to ISA hierarchies? What other extensions of association rules are you farniliar with? (Sections 26.3.3 and 26.3.4)
- What is a sequential pattern? How can we cornpute sequential patterns? (Section 26.3.5)
- Can we use association rules for prediction? (Section 26.3.6)
- What is the difference bet\'leen Bayesian Networks and association rules? (Section 26.3.7)
- Can you give exanlples of classification and regression rules? How is support and confidence for such rules defined? (Section 26.3.8)
- What are the cOlnponents of a decision tree? How are decision trees constructed? (Sections 26.4.1 and 26.4.2)
- What is a cluster? What inforrnation do we usually output for a cluster? (Section 26.5)
- How can we define the distance between two sequences? Describe an algorithnl to find all sequences similar to a query sequence. (Section 26.6)
- Describe the block evolution Inodel and define the problcllS of increlnental rnodel maintenance and change detection. What is the added challenge in rnining data strearns? (Section 26.7)
- Describe an incrernental algorithnl for conlpllting frequent iternsets. (Section 26.7.1)
- Give exarnples of other tasks related to data rnining. (Section 26.8)

EXERCISES

Exercise 26.1 Briefly ans\ver the following questions:

1. Define *support* and *confidence* for an association 1"ule.
2. Expla.in why association rules cannot be used directly for prediction, \vithout further analysis or clornain knowledge.
3. What are the differences between *association rules*, *classification rules*, and *regression rules*?
4. \What is the difference between *classification* and *clustering*?

111	201	5/1/2002	ink	1
111	201	5/1/2002	lnilk	2
11.1	201	5/1/2002	.JuIce	1
112	105	6/3/2002	pen	1
112	105	6/3/2002	ink	1
112	105	6/3/2002	water	1
113	106	5/10/2002	pen	1
113	106	5/10/2002	water	2
113	106	5/10/2002	milk	1
114	201	6/1/2002	pen	2
114	201	6/1/2002	ink	2
114	201	6/1/2002	JUIce	4
114	201	6/1/2002	water	1
114	201	6/1/2002	ruilk	1

Figure 26.16 The Purchases2 Relation

5. What is the role of information visualization in data mining?
6. Give exarrlples of queries over a database of stock price quotes, stored as sequences, one per stock, that cannot be expressed in SQL.

Exercise 26.2 Consider the Purchases table shown in Figure 26.1.

1. Simulate the algorithrn for finding frequent iterllsets on the table in Figure 26.1 with *minsup=90* percent, and then find association rules with *m,inconJ=90* percent.
2. Can you modify the table so that the same frequent itemsets are obtained with *'fninsup=90* percent as with *minsup*=70 percent on the table shown in Figure 26.1?
3. Sirllulate the algorithrIl for finding frequent iternsets on the table in Figure 26.1 with *rn'insup=lO* percent and then find association rules with *rninconj=90* percent.
4. Can you modify the table so that the sarne frequent iternsets are obtained with *minsup*=10 percent as with *minsup*=70 percent on the table shown in Figure 26.1?

Exercise 26.3 Assulne we are given a dataset D of rnarket baskets and have computed the set of frequent iternsets $\mathcal{X}$ in 1) for a given support threshold *minsup*. Assume that we would like to add. another dataset D' to D, and rnaintain the set of frequent itmnsets with support threshold ***minsup*** in D U D'. Consider the following algorithrIl for incrernental Inaintenance of a set of frequent iternsets:

1. We run the ***a priori*** algoritlun on D' and find all frequent iterllsets in D' and their support. The result is a set of iterllsets $\mathcal{X}'$. We also cornpute the support of all itcrnsets X E $\mathcal{X}$ in $J)'$.
2. We then rnake a scan over D to cornpute the support of all iternsets in $\mathcal{X}'$.

Answer the following questions about the algorithm:

- The last step of the algorithm is rnissing; that is, what should the algorithm output'?
- Is this algorithm lllore efficient than the algorithm described in Section 26.7.1'1

Exercise 26.4 Consider the Purchases2 table shown in Figure 26.16.

- List all iterllsets in the negative border of the dataset.
- List all frequent itelnsets for a support threshold of 50%.
- Give an exaruple of a database in which the addition of this database does not change the negative border.
- Give an exarnple of a database in which the addition of this database would change the negative border.

Exercise 26.5 Consider the Purchases table shown in Figure 26.1. Find all (generalized) association rules that indicate the likelihood of items being purchased on the same date by the same customer, with *minsup* set to 10% and *minconj* set to 70%.

Exercise 26.6 Let us develop a new algorithm for the computation of all large itemsets. Assume that we are given a relation D siInilar to the Purchases table shown in Figure 26.1. We partition the table horizontally into k parts $D_1, \ldots, D_k$.

1. Show that, if itemset X is frequent in D, then it is frequent in at least one of the k parts.
2. Use this observation to develop an algorithm that cornputes all frequent itemsets in two scans over D. (Hint: In the first scan, compute the locally frequent itemsets for each part D_i, i E {I, ... , k}.)
3. Illustrate your algorithm using the Purchases table shown in Figure 26.1. The first partition consists of the two transactions with *transid* 111 and 112, the second partition consists of the two transactions with *transid* 113 and 114. Assulne that the minimum support is 70 percent.

Exercise 26.7 Consider the Purchases table shown in Figure 26.1. Find all sequential patterns with *minsup* set to 60%. (The text only sketches the algorithm for discovering sequential patterns, so use brute force or read one of the references for a complete algorithm.)

Exercise 26.8 Consider the SubscriberInfo Relation shown in Figure 26.17. It contains information about the marketing cmnpaign of the *DB Aficionado* magazine. The first two colurnns show the age and salary of a potential customer and the *subscription* colurnn shows whether the person subscribes to the rnagazine. \Ve want to use this data to construct a decision tree that helps predict whether a person will subscribe to the 11lagazine.

1. Construct the AVC-group of the root node of the tree.
2. Assume that the spliting predicate at the root node is $age \leq 50$. Construct the AVC-groups of the two children nodes of the root node.

Exercise 26.9 Assurne you are given the following set of six records: (7,55), (21, 202), (25,220), $\langle 12, 73\rangle$, $\langle 8, 61\rangle$, and $\langle 22, 249\rangle$.

1. Assurning that all six records belong to a single cluster, cornpute its center and radius.
2. Assurne that the first three records belong to one cluster and the second three records belong to a different cluster. COlnpute the center and radius of the two clusters.
3. \Vhich of the two clusterings is 'better' in your opinion and why?

Exercise 26.10 Asslune you are given the three sequences $(1, 3, 4)$, $\langle 2, 3, 2\rangle$, $\langle 3, 3, 7\rangle$. COlnpute the Euclidian Bonn between all pairs of sequences.

age	salary	subscription
37	45k	No
39	70k	Yes
56	50k	Yes
52	43k	Yes
35	90k	Yes
32	54k	No
40	58k	No
55	85k	Yes
43	68k	Yes

Figure 26.17 The SubscriberInfo Relation

BIBLIOGRAPHIC NOTES

Discovering useful knowledge from a large database is lllore than just applying a collection of data rnining algorithms, and the point of view that it is an iterative process guided by an analyst is stressed in [265] and [666]. Work on exploratory data analysis in statistics, for example [745], and on rnachine learning and knowledge discovery in artificial intelligence was a precursor to the current focus on data lTlining; the added ernphasis on large volunles of data is the inlportant new elernent. Good recent surveys of data mining algorithms include [267, 397, 507]. [266] contains additional surveys and articles on many aspects of data mining and knowledge discovery, including a tutorial on Bayesian networks [:371]. The book by Piatetsky-Shapiro and Frawley [595] contains an interesting collection of data rnining papers. The annual SIGKDD conference, run by the ACM special interest group in knowledge discovery in databases, is a good resource for readers interested in current research in data mining (25, 162, 268, 372, 613, 691], as is the *Journal of Knowledge D'iscovery and Data Mining.* [363, 370, 511, 781] are good, in-depth textbooks on data nlining.

The problern of mining association rules was introduced by Agrawal, Itnielinski, and Swami [20]. lVlany efficient algorithnls have been proposed for the cornputation of large iternsets, including [21,117,364,683,738,786].

Iceberg queries have been introduced by F'ang et al. [264]. There is also a large body of research on generalized forrns of <lssociation rules; for example, [700, 701, 703]. The problem of finding rnaxirnal frequent itelnsets has also received significant attention [13, 67, 126, 346, 347, 479, 787]. Algorithrns for rnining association rules with constraints are considered in [68,462, 563, 590, 591, 703].

Parallel algorithnls are described in [23] and [655]. Recent papers on parallel data ruining can be found in [788], and work on distributed data 111ining can be found in [417].

[291] presents an aigoritlull for discovering association rules over a continuous nUllwric attribute; association rules over numeric attributes are also discussed in [78:3J. The general fornl of association rules, in which attributes other than the transaction id are grouped is developed in [529]. Association rules over iterns in a hierarchy are discllssed in [361, 700]. Further

extensions and generalization of association rules are proposed in [67, 115, 563]. Integration of rnining for frequent itemsets into database systcrns has been addressed in [654, 743]. The problern of Inining sequential patterns is discussed in [24], and further algorithrIls for rnining sequential patterns can be found in [510, 702].

General introductions to classification and regression rules can be found in [362, 532]. The classic reference for decision and regression tree construction is the CART book by Breilnan, Friedmau, Olsheu, and Stone [111]. A lnachine learning perspective of decision tree construction is given by Quinlan [603]. Recently, several scalable algorithnls for decision tree construction have been developed [309, 311, 521, 619, 674J.

'rhe clustering problern has been studied for decades in several disciplines. Sample textbooks include [232, 407, 418]. Scalable clustering algorithuls include CLARANS [562], DBSCAN [249, 250], BIRCH [798], and CURE [344]. Bradley, Fayyad, and Reina address the problem of scaling the K-Means clustering algorithm to large databases [108, 109]. The problern of finding clusters in subsets of the fields is addressed in [19]. Ganti et al. exauline the problerll of clustering data in arbitrary rnetric spaces [302]. Algorithrlls for clustering caterogical data include STIRR [315J and CACTUS [301]. [651] is a clustering algorithm for spatial data.

Finding siulilar sequences from a large database of sequences is discussed in [22, 262, 446, 606,680].

Work on incrernental rnaintenance of association rules is considered in [174, 175, 736]. Ester et al. describe how to nlaintain clusters incrernentally [248], and Hidber describes how to rnaintain large iteulsets incrernentally [378]. There has also been recent work on rnining data streams, such as the construction of decision trees over data streams [228, 309, 393] and clustering data streanlS [343,568]. A general framework for ruining evolving data is presented in [299]. A framework for measuring change in data characteristics is proposed in [300J.

27

INFORMATION RETRIEVAL ANDXMLDATA

- How are DBMSs evolving in response to the growing aIllounts of text data?
- What is the vector space rnodel and how does it support text search?
- How are text collections indexed?
- Cornpared to IR systenls, what is new in Web search?
- How is XML data different from plain text and relational tables?
- What are the main features of XQuery?
- What are the irnplementation challenges posed by XML data?
- Key concepts: information retrieval, boolean and ranked queries; relevance, precision, recall; vector space model, TF/IDF terrn weighting, document similarity; inverted index, signature file; Web crawler, hubs and authorities, Pigeon Rank of a webpage; sernistructured data lllodel, XML; XQuery, path expressions, FLWR queries; XML storage and indexing

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A *memex* is a device in which an individual stores all his books, records, and cornrnunications, and which is rnechanized so that it rnay be consulted with exceeding speed and flexibility.

—Vannevar Bush, *As We May Think, 1945*

The field of inforlnation retrieval (IR) has studied the problenl of sea,rching collections of text docurnents since the 19508 and developed largely independently of database systenls. The proliferation of text docunlents on the Web lllade docurnent search an everyday operation for 1110st people and led to renewed research on the topic.

The database field's desire to expand the kinds of data that can be managed in a DBMS is well-established and reflected in developments like object-relational extensions (Chapter 23). Documents on the Web represent one of the rnost rapidly growing sources of data, and the challenge of rnanaging such documents in a DBMS has naturally become a focal point for database research.

The Web, therefore, brought the two fields of database rnanagement systenls and information retrieval closer together than ever before, and, as we will see, XML sits squarely in the middle ground between thenl. We introduce IR systems as well as a data model and query language for XML data and discuss the relationship with (object-)relational database systerns.

In this chapter, we present an overview of information retrieval, Web search, and the emerging XML data model and query language standards. We begin in Section 27.1 with a discussion of how these text-oriented trends fit within the context of current object-relational database systeIns. We introduce inforrnation retrieval concepts in Section 27.2 and discuss specialized indexing techniques for text in Section 27.3. We discuss Web search engines in Section 27.4. In Section 27.5, we briefly outline current trends in extending database systems to support text data and identify SOllle of the irnportant issues involved. In Section 27.6, we present the XML data lllodel, building on the XML concepts introduced in Chapter 7. We describe the XQuery language in Section 27.7. In Section 27.8, we consider efficient evaluation of XQuery queries.

27.1 COLLIDING WORLDS: DATABASES, IR, AND XML

'The \i*Teb* is the rnost widely used doculnent collection today, and search on the Web differs froIn traditional IR-style docurnent retrieval in iluportant ways. First, there is great emphasis on scalability to very large document collections. IR systerns typically dealt with tens of thousands of documents, whereas the Web contains billions of pages.

Second, the Web has significantly changed how docurnent collections are created and used. Traditionally, IR systerlls were aimed at professionals like librarians and legal researchers, who were trained in using sophisticated retrieval engines. Docurnents were carefully prepared, and docllrnents in a given collection were typically on related topics. On thevVeb, docurnents are created by an infinite

variety of individuals for equally many purposes, and reflect this diversity in size and content. Searches are carried out by ordinary people with no training in using retrieval software.

The ernergence of XML has added a third interesting diInensioI1 to text search: Every cloClunent can no\v be rnarked up to reflect additional infol'lnation of interest, such as authorship, source, and even details about the intrinsic content. This has changed the nature of a "document" froIn free text to textual objects with associated fields containing metadata (data about data) or descriptive infonnation. Links to other docurnents are a particularly irnportant kind of lnetadata, and they can have great value in searching docurnent collections on the Web.

The Web also changed the notion of what constitutes a docunlent. Documents on the Web may be multinledia objects such as irnages or video clips, with text appearing only in descriptive tags. We must be able to Inanage such heterogeneous data collections and support searches over thern.

Database rnanagernent systenls traditionally dealt with simple tabular data. In recent years, object-relational database systerns (ORDBMSs) were designed to support complex data types. Images, videos, and textual objects have been explicitly rnentioned as exaruples of the data types ORDBMSs are intended to support. Nonetheless, current database systerns have a long way to go before they can support such cOlnplex data types satisfactorily. In the context of text and XML data, challenges include efficient support for searches over textual content and support for searches that exploit the loose structure of XML data.

27.1.1 DBMS versus IR Systems

Database and IR systcrns have the COllllnon objective of supporting searches over collections of data. However, rnany irnportant differences have influenced their developrnent.

- Searches versus Queries: IR systerns are designed to support a specialized class of queries that we also call searches. Searches are specified in ternlS of a few search terms, and the underlying data is usually a collection of unstructured text docurnents. III addition, an irnportant feature of IR searches is that search resultsrnay be ranked, or ordered, in tcrrns of how 'well' the search results rnatch the search terms. In contrast, database systerns support a very general class of queries, and the underlying data is rigidly structured. Unlike III systems, database systerns have traditionally returnedunranked sets of results. (Even the recent SQL/OLAP extensions that support early results and searches over ordered data (see Chapter 25)

do not order results in terlTIS of how well they rnatch the query. Relational queries are *precise* in that a ro\v is either in the answer or it is not ; there is no notion of 'how well a row matches' the query.) In other words, a relational query only assigns two ranks to a row, indicating 'whether the row is in the ans\ver or not.

- Updates and Tr'ansactions: IR systelns are optirnized for a read-Illostly workload and do not support the notion of a transaction. In traditional IR systerlls, ne\v docurnents are added to the doculnent collection frorH tirne to time, and index structures that speed up searches are periodically rebuilt or updated. Therefore, docllrnents that are highly relevant for a search rnight exist in the IR systeln, but not be retrievable yet because of outdated index structures. In contrast, database systerns are designed to handle a wide range of workloads, including update-intensive transaction processing workloads.

These differences in design objectives have led, not surprisingly, to very different research elnphases and system designs. Ilesearch in IR studied ranking functions extensively. For example, arllong other topics, research in IR investigated how to incorporate feedback frOITl a user's behavior to modify a ranking function and how to apply linguistic processing techniques to improve searches. Database research concentrated on query processing, concurrency control and recovery, and other topics, as covered in this book.

The differences between a DBMS and an IR systenl from a design and irnplementation standpoint should become clear as we introduce IR systerlls in the next few sections.

27.2 INTRODUCTION TO INFORMATION RETRIEVAL

There are two COrllr110n types of searches, or queries, over text collections: boolean queries and ranked queries. In a boolean **query**, the user specifies an expression constructed using terlllS and boolean operators (And, Or, Not). For exalnple,

database And *(lvlicT08ojt* Or *IBM)*

This query asks for all docurnents that contain the terrn *database* and in addition, either *Microsoft* or *IBM.*

In a ranked query the user specifies one or rnore terrlls, and the result of the query is a list of docurllents ranked by their relevance to the query. Intuitively, docurnents at the top of the result list are expected to 'rnatch' the search

docid	Document
1	agent Janles Bond good agent
2	agent rnobile cornputer
3	James Madison Inovie
4	Janles Bond movie

Figure 27.1 A Text Database with Four Records

condition ruore closely, or be 'rnore relevant', than doculnents lower in the result list. While a document that contains *Microsoft* satisfies the search'*Microsoft, IBM,'* a document that also contains *IBM* is considered to be a better match. Similarly, a docunlent that contains several occurrences of *Microsoft* might be a better rnatch than a document that contains a single occurence. Ranking the docurnents that satisfy the boolean search condition is an important aspect of an IR search engine, and we discuss how this is done in Sections 27.2.3 and 27.4.2.

An important extension of ranked queries is to ask for documents that are most relevant to a given natural language sentence. Since a sentence has linguistic structure (e.g., subject-verb-object relationships), it provides more information than just the list of words that it contains. We do not discuss natural language search.

27.2.1 Vector Space Model

We now describe a widely-used franlework for representing docurnents and searching over docurnent collections. Consider the set of all terrns that appear in a given collection of documents. We can represent each document as a vector with one entry per ternl. In the shnplest 1'01'111 of doclunent vectors, if terrn *j* appears *k* tirnes in dOCUlnent i, the document vector for docurnent i contains value *k* in position *j*. The docurnent vector for *i* contains the value 0 in positions corresponding to terrns that do not appear in i.

Consider the exaInple collection of four docurnents shown in Figure 27.1. rrhe docUluent vector representation is illustrated in Figure 27.2; each row represents a docurnent. This representation of docurnents as terrn vectors is called the vector space model.

docid	agent	Bond	computer	good	James	Madison	mobile	movie
1	2	1	0	1	1	0	0	0
2	1	0	1	0	0	0	1	0
3	0	0	0	0	1	1	0	1
4	0	1	0	0	1	0	0	1

Figure 27.2 Document Vectors for the Example Collection

27.2.2 TFIIDF Weighting of Terms

We described the value for a terril in a document vector as simply the term frequency (TF), or nurnber of occurrences of that terrn in the given document. This reflects the intuition that a term which appears often is more iInportant in characterizing the document than a terrn that appears only once (or a term that does not appear at all).

However, some terms appear very frequently in the document collection, and others are relatively rare. The frequency of terms is eITlpirically observed to follow a Zipfian distribution, as illustrated in Figure 27.3. In this figure, each position on the X-axis corresponds to a terrll and the Y-axis corresponds to the nUlnber of occurrences of the term. Terms are arranged on the X-axis in decreasing order by the nurnber of tirnes they occur (in the docurnent collection as a whole).

As rnight be expected, it turns out that extremely COmlTIOn terms are not very useful in searches. Examples of such common terms include *a*, *an*, *the* etc. Terrns that occur extremely often are called stop words, and docunlents are pre-processed to elirilinate stop words.

Even after eliminating stop words, we have the phenorilenon that some words appear nluch luore often than others in the docurnent collection. Consider the words *Linux* and *kernel* in the context of a collection of dOCUlnents about the Linux operating systern. While neither is COlnrnon enough to be a stop word, *Linux* is likely to appear much rnore often. Given a search that contains both these keywords, we are likely to get better results if we give Inore irnportance to docurnents that contain *kernel* than docurnents that contain *Linux*.

We can capture this intuition by refining the docurnent vector representatioll as follows. The value associated with ternl j in the docurnent vector for docurnent i, denoted as w_{ij}, is obtained by rnultiplying the terrIl frequency tij (the nuruber of tirnes term j appears in docurnent i) by the inverse docurnent frequency (IDF) of terrn j in the docurnent collection. IDF of a tenn j is defined as

log(lVInj); where N is the total rHnnber of dOCUlnents, and n_j is the nurnber of cloClllnents that tenn *j* appears in. This effectively increases the weight given to rare tenns. As an example, in a collection of 10,000 docurnents, a terrIl that appears in half the docurnents has an IDF of 0.3, and a tenll that occurs in just one docurnent has an IDF of 4.

Length Normalization

Consider a docurnent *!J.* Suppose that we lnodify it by adding a large nUlllber of new terrns. Should a the weight of a terrn *t* that appears in *D* be the saIne in the doclunent vectors for *D* and the rnodified dOCUlTlent? Although the TFjIDF weight for t is indeed the saIne in the two document vector, our intuition suggests that the weight should be less in the 1110dified document. Longer docul'llents tend to have lnore terms, and lnore occurrences of any given terrn. Thus, if two doculnents contain the saIne nUlnber of occurrences of a given tenll, the importance of the ten'll in characterizing the document also depends on the length of the docull1ent.

Several approaches to length nornlalization have been proposed. Intuitively, all of ther'll reduce the irnportance given to how often a term occurs as the frequency grows. In traditional IR systelns, a popular way to refine the sirnilarity Inetric is cosine length normalization:

$$w_{ij}^{*} = \frac{w_{ij}}{\sqrt{\sum_{k=1}^{t} w_{ik}^2}}$$

In this formula, t is the nurnbei' of tenns in the dOCulnent collection, w_{ij} is the TFjIDF weight without length norrnalization, and w_{ij}^{*} is the length adjusted TFjlDF weight.

Tenns that occur frequently in a doculnent are particularly problenlatic on the Web because webpages are often deliberately rnodified by adding rnany copies of certain words···· for exarnple, sale, free, sex to increase the likelihood of their being returned in response to queries. For this reason, Web search engines typically norrnalize for length by imposing a lnaxirnurn value (usually 2 or 3) for terrIl frequencies.

27.2.3 Ranking Document Similarity

We now consider how the vector space representation allows us to rank dOCIIrnents in the result of a ranked query. A key observation is that a ranked query can itself be thought of as a docUlllent, since it is just a collection of terrl1s. 1'his allows us to use document similarity as the basis for ranking query

results---the doculnent that is rnost sirnilar to the query is ranked highest, and the one that is least sirnilar is ranked lowest.

If a total of t teru18 appear in the collection of docurnents (t is 8 in the exaulple shown in Figure 27.2), we can visualize document vectors in a t-diInensional space in \vhich each axis is labeled with a te1'1n. This is illustrated in Figure 27.4, for a two-dirnensional space. The figure shows doculuent vectors for two documents, D_1 and $.D_2$, as well as a query Q.

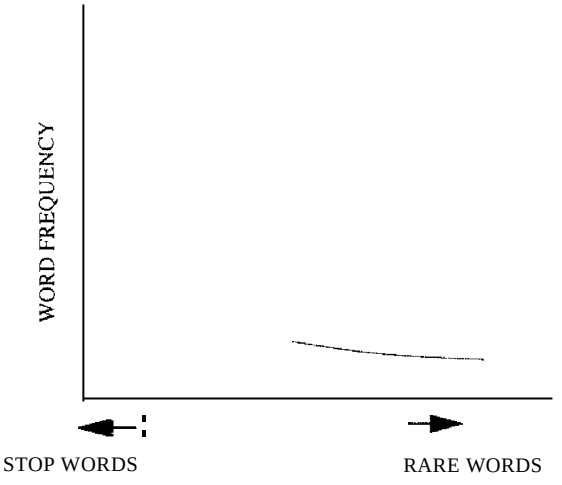


Figure 27.3 Zipfian Distribution of Term Frequencies

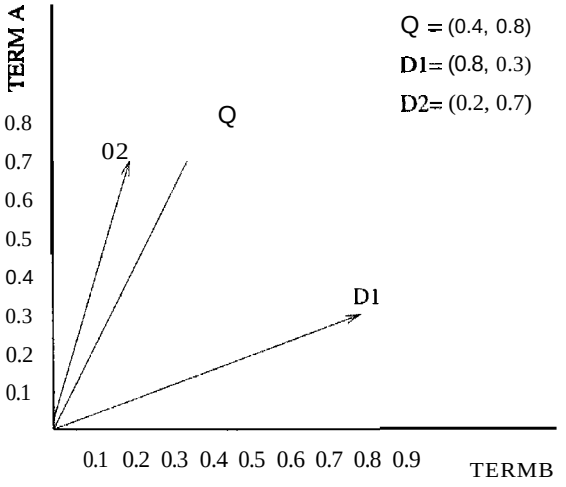


Figure **27.4** Document Similarity

The traditional rneasure of closeness between two vectors, their *dot product*, is used as a lneasure of docurnent siInilarity. The siInilarity of query Q to a doculnent D_i is Illea8Ured by their dot produet:

$$sim(Q, D_i) \text{ —} \sum_{.1=1}^{t} q_j^* . w_{ij}^*$$

In the example shown in Figure 27.4, $sim(Q, D_1)$ $(0.4 * 0.8)$.+ $(0.8 * 0.3)$ = 0.56, and *si1n(Q, D2)* = $(0.4 * 0.2)$-+- $(0.8 * 0.7)$ = 0.64. Accordingly, *D2* is ranked higher than 1)1 in the search result.

In the context of the Web, docurnent sirnila1'ity is one of several IneaSU1'es that can be used. to rank results, but should not be used exclusively. First, it is questionable whether users want dOCllrnents that are sirnilar to the query (which typically consists of one or two 'words) or dOCUll.lenS that contain useful inforrnation related to the quer,Y te1'111S. IntuitivelY,we want to give ilnportance to the *quality* of a Web page while ranking it, in addition to reflecting the sirnilarity of the page to a given query. Links between pages provide valuable

additional inforrnation that can be used to obtain high-quality results. We discuss this issue in Section 27.4.2.

27.2.4 Measuring Success: Precision and Recall

Two criteria are cornnlonly used to evaluate information retrieval systerlls. Precision is the percentage of retrieved documents that are relevant to the query. Recall is the percentage of relevant docurnents in the database that are retrieved in response to a query.

Retrieving all documents in response to a query trivially guarantees perfect recall, but results in very poor precision. The challenge is to achieve good recall together with high precision.

In the context of search over the Web, the size of the underlying collection is on the order of billions of docuruents. Given this, it is questionable whether the traditional measure of recall is very useful. Since users typically don't look beyond the first screen of results, the quality of a Web search engine is largely deterlnined by the results shown on the first page. The following adapted definitions of precision and recall rnight be more appropriate for Web search engInes:

- Web Search Precision: The percentage of results on the first page that are relevant to the query.
- Web Search Recall: rrhe fraction N/M, expressed as a percentage, where M is the nUluber of results displayed on the front page, and of the M ruost relevant documents, N is the number displayed on the front page.

27.3 INDEXING FOR TEXT SEARCH

In this section, we introduce two indexing techniques that support the evaluation of boolean and ranked queries. 'The *'inverted index* structure discussed in Section 27.3.1 is widely used due to its sirnplicity and good perforlnance. Its rnain disadvantage is that it imposes a significant space overhead: The size can be up to 300 percent the size of the original file. The *signature file* index discussed in Section 27.3.2 has a sInall space overhead and offers a quick filter that elirninates rnost nonqualifying docurnents. However, does not scale as well to larger datahase sizes because the index has to be sequentially scanned.

Before a doeuruent is indexed, it is typically pre-processed to elirninate stop words. Since the size of the indexes is very sensitive to the nurnber of tern1S in the docurnent collection, elirninating stop words can greatly reduce index

size. IR systcrns also do certain other kinds of pre-processing. :For instance, they apply stelnming to reduce related terrns to a ca,nonical forrn. This step also reduces the nUluber of terrn8 to be indexed, but equally irnportantly, it allows us to retrieve documents that lnay not contain the exact query terrIl but contain S(Hne variant. As an example, the terrns *run*, *running*, and *runner* all stern to *run*. The terrIl *run* is indexed, and every occurrence of a variant of this term is treated as an occurrence of *run*. A query that specifies *runner* finds docurnents that contain any word that stenlS to *run*.

27.3.1 Inverted Indexes

An inverted index is a data structure that enables fast retrieval of all documents that contain a query terr11. For each ternl, the index rnaintains a list (called the inverted list) of entries describing occurrences of the tenn, with one entry per docurnent that contains the ternl.

Consider the inverted index for our running example shown in Figure 27.5. The term 'Jarnes' has an inverted list with one entry each for documents 1, 3, and 4; the term 'agent' has entries for docurnents 1 and 2.

The entry for document *d* in the inverted list for terrn t contains details about the occurrences of term *t* in document *d*. In Figure 27.5, this information consists of a list of locations within the document that contain term *t*. Thus, the entry for document 1 in the inverted list for terrn 'agent' lists the locations 1 and 5, since 'agent' is the first and fifth word of docurnent 1. In general, we can store additional information about each occurrence (e.g., in an HTML docurnent, is the occurrence in the TITLE tag?) in the inverted list. We can also store the length of the docurnent if this is used for length norlnalization (see below).

The collection of inverted lists is called the postings file. Inverted lists can be very large for large doeurnent collections. In fact, Web search engines typically store each inverted list on a separate page, and Inost lists span rnultiple pages (and if so, are rnaintained as a linked list of pages). In order to quickly find the inverted list for a, query terrn, all possible query terrns are organized in a second index structure such as a B+ tree or a hash index.

The second index, called the lexicon, is Inuch srnaller than the postings file since it only contains one entry per terrn, and further, only contains entries for the set of terl11S that are retained after elirninating stop words, and applying stenlluing rules. An entry consists of the terlIl, SOH1C surnrnary inforrnation about its inverted list, and the address (on disk) of the inverted list. In Figure 27.5, the SllIl1lnary inforrnation consists of the lllllllber of entries in the inverted

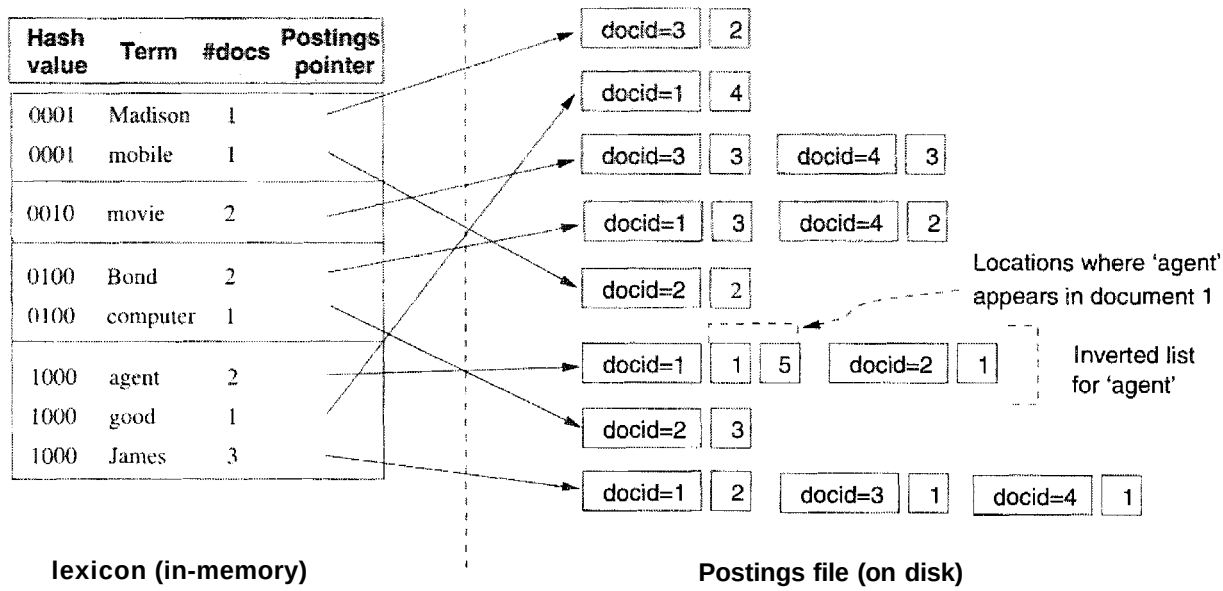


Figure 27.5 Inverted Index for Example Collection

list (i.e., the nurnber of documents that the terl11 appears in). In general, it could contain additional infonnation such as the IDF for the terrIl, but it is il11portant to keep the entry's size as slllall as possible.

The lexicon is rua.intained in-lllelnory, and enables fast retrieval of the inverted list for a query terrn. rrhe lexicon in Figure 27.5 uses a hash index, and is sketched by showing the hash value for the terrn; entries for terms are grouped into hash buckets by their hash value.

Using an Inverted Index

A_ query containing a single tenn is evaluated by first searching the lexicon to find the address of the inverted list for the terrIl. Then the inverted list is retrieved, the docids in it are rnapped to physical doculnent addresses, and the corresponding docurnents are retrieved. If the results are to be ranked, the relevance of each docurnent in the inverted list to the query term is C01l1puted, and docurnents are then retrieved in order of their relevance rank. ()bserve that the inforrna,tion needed to cornpute the relevance measure described in Section 27.2 -- the frequency of the query ternl in the dOCu1nent, the IDF of the terrn in the docurnent collection, and the length of the docurnent if it is used for length nonnalizatioll-------are all available in either the lexicon or the inverted list.

When inverted lists are very long, as in Web search engines, it is useful to consider \vhether we should precornpute the relevance of each dOCUlnent in the inverted list for a terrn (with respect to that terrn) and sort the list by relevance rather than docurnent id. This would speed up querying because we can just

look at a prefix of the inverted list, since users rarely look at 11101'0 than the first few results. However, maintaining lists in sorted order by relevance can be expensive. (Sorting by dOcUll1cnt id is convenient because new dOCUlnents are assigned increasing ids, and we can therefore sirnply append entries for new dOCUlnents at the end of the inverted list. Further, if the sirnilarity function is changed, we do not have to rebuild the index.)

A query with a conjunction of several terrns is evaluated by retrieving the inverted lists of the query terrns one at a time and intersecting theln. In order to rninirnize memory usage, the inverted lists should be retrieved in order of increasing length. A query with a disjunction of several terrns is evaluated by ruerging all relevant inverted lists.

Consider the exaruple inverted index shown in Figure 27.5. To evaluate the query 'JaUles', we probe the lexicon to find the address of the inverted list for 'Ja1nes', fetch it from disk and then retrieve docurIlent 1. To evaluate the query ,Jarnes' AND 'Bond', we first retrieve the inverted list for the tenn 'Bond' and intersect it with the inverted list for the terrn 'Janles.' (The inverted list of the terrn 'Bond' has length two, whereas the inverted list of the terrIl 'Jarnes' has length three.) rrhe result of the intersection of the list (1,4) with the list (1, 3, 4) is the list (1,4) and doculuents 1 and 4 are therefore retrieved. 'fa evaluate the query '.lalnes' OR 'Bond,' we retrieve the two inverted lists in any order and merge the results.

For ranked queries with lnultiple tenns, we 1nust fetch the inverted lists for all terrl1s, COlllpute the relevance of every doclunent that appears in one of these lists with respect to the given collection of query terrns, and then sort the docurnent ids by their relevance before fetching the docluuents in relevance rank order. Again, if the inverted lists are sorted by the relevance measure, we can support ranked queries by typically processing only sluall prefixes of the the inverted lists. (()bserve that the relevance of a docuIllent with respect to the query is easily cornputed froIn its relevance with respect to each query term.)

27.3.2 Signature Files

A signature file is another index structure for text database systerns that supports efficient evaluation of boolean queries. A signature file contains an index record for each docurnent in the database. This index record is called the signature of the dOClunent. Each signature has a fixed size of b bits; b is called the signature width. rrhe bits that are set depend on the words that appear in the docllrnent. We rnap words to bits by applying a hash function to ea,ch word in the docurnent and we set the bits that appear in the result of

1	agent Jarnes Bond good agent	1100
2	agent mobile C0111puter	1101
3	James Madison lnovie	1011
4	JaInes Bond rnovie	1110

Figure 27.6 Signature File for Example Collection

the hash function. Note that unless we have a bit for each possible word in the vocabulary, the same bit could be set twice by different words because the hash function maps both words to the saIne bit. We say that a signature S_1 matches another signature 82 if all the bits that are set in signature 8_2 are also set in signature 8_1. If signature 8_1 Inatches signature 8_2, then signature 8_1 has at least as many bits set *as* signature 8_2.

For a query consisting of a conjunction of terms, we first generate the query signature by applying the hash function to each word in the query. We then scan the signature file and retrieve all documents whose signatures match the query signature, because every such document is a potential result to the query. Since the signature does not uniquely identify the words that a docuInent contains, we have to retrieve each potential rnatch and check whether the docunlent actually contains the query terms. A docurnent whose signature matches the query signature but that does not contain all terms in the query is called a false positive. A false positive is an expensive rnistake since the docurnent has to be retrieved froln disk, parsed, stemrned, and checked to determine whether it contains the query terms.

For a query consisting of a disjunction of tenns, we generate a list of query signatures, one for each terrn in the query. The query is evaluated by scanning the signature file to find docurnents whose signatures rnatch any signature in the list of query signatures.

As an exarllple, consider the signature file of width 4 for our running exarnple shown in Figure 27.6. rrhe bits set by the hashed values of all query terrns are shown in the figure. To evaluate the query 'Jal11es,' we first cOlnpute the hash value of the terrn; this is 1000. Then we scan the signature file and find rnatching index records. As we can see fronl Figure 27.6, the signatures of all records have the first bit set. We retrieve all doculnents and check for false positives; the only false positive for this query is docurnent with rid 2. (lJnfortunately, the hashed value of the terrn 'agent' also happened to set the very first bit in the signature.) Consider the query 'James' And 'Bond.' The query signature is llOO and three docurnent signatures rnatch the query signature. Again, \ve retrieve one false positive. As another exarnple of a conjunctive query, con-

sider the query 'movie' And 'Madison.' The query signature is 0011, and only one doclunent signature lnatches the query signature.No false positives are retrieved.

Note that for each query we have to scan the cOlllplete signature file, and there are as 11lany records in the signature file as there are documents in the database. To reduce the anlount of data that has to be retrieved for each query, we can vertically partition a signature file into a set of bit slices, and we call such an index a bit-sliced signature file. The length of each bit slice is still equal to the number of doculllents in the database, but for a query with q bits set in the query signature we need only to retrieve q bit slices. The reader is invited to construct a bit-sliced signature file and to evaluate the exarnple queries in this paragraph using the bit slices.

27.4 WEB SEARCH ENGINES

Web search engines rIlust contend with extreruely large nurubers of doculllents, and have to be highly scalable. Docurnents are also linked to each other, and this link infonnation turns out to be very valuable in finding pages relevant to a given search. These factors have caused search engines to differ frorn traditional IR systerns in irnportant ways. Nonetheless, they rely on sorne forn1 of inverted indexes as the basic indexing mechanism. In this section, we discuss Web search engines, using Google as a typical example.

27.4.1 Search Engine Architecture

Web search engines crawl the web to collect docurnents to index. 'Ihe crawling algorithrn is sirrlple, but crawler software can be cornplex because of the details of connecting to millions of sites, minimizing network latencies, parallelizing the crawling, dealing with tirneouts and other connection failures, ensuring that crawled sites are not unduly stressed by the cra\vler, and other practical concerns.

The search algorithrn used by a crawler is a graph traversal. Starting at a collection of pages with rnany links (e.g.,Yahoo directory pages), all links on cra\vled pages are followed to identify new pages. This step is iterated, keeping track of which pages have been visited in order to avoid re-visiting thenl.

The collection of pages retrieved through crawling can be enonnous, on the order of billions of pages. Indexing thern is a very expensive task. Fortunately, tlle task is highly parallelizable: Each docurnent is independently arlalyzed to create inverted lists for the terrns that appear in the docurnent. These per-doCUlnent lists are then sorted by terrn and luerged to create cornplete per-

term inverted lists that span all dOCllrnents. Ternl statistics such as IDF can be cornputed during the lnerge phase.

Supporting searches over such vast indexes is another luanulloth undertaking. Fortunately, again, the task is readily parallelized using a cluster of inexpensive Inachines: We can deal with the anlount of data by partitioning the index across several rnachines. Each Inachine contains the inverted index for those terms that are Inapped to that luachine (e.g., by hashing the tenn). Queries 111ay have to be sent to luultiple Inachines if the terrl1S they contain are handled by different rnachines, but given that Web queries rarely contain rnore than two terrns, this is not a serious probleln in practice.

We rnust also deal \vith a huge volume of queries; Google supports over 150 lllillion searches each day, and the nUlnber is growing. This is acc(nnplished by replicating the data across several machines. We already described how the data is partitioned across Inachines. For each partition, we now assign several nlachines, each of which contains an exact copy of the data for that partition. Queries on this partition can be handled by any rnachine in the partition. Queries can be distributed across rnachines on the basis of load, by hashing on IP addresses, etc. Replication also addresses the problern of high-availability, since the failure of a Inachine only increases the load on the remaining rnachines in the partition, and if partitions contain several rnachines the ilnpact is sIuall. Failures can be rnade transparent to users by routing queries to other Inachines through the load balancer.

27.4.2 Using Link Information

webpages are created by a variety of users for a variety of purposes, and their content does not always lend itself to effective retrieval. The rnost relevant pages for a search rnay not contain the search terrns at all and are therefore not returned by a boolean keyword search! For exarnple, consider the query ternl ‘Web browser.' A boolean text query using the tenns does not return the relevant pages of Netscape Corporation or Microsoft, because these pages do not contain the terrn ‘Web browser' at all. Sirnilarly, the horne page of 'Yahoo does not contain the terrn 'search engine.' The problenl is that relevant sites do not necessarily describe their contents in a way that is llseful for boolean text queries.

Until now, we only considered infonnation 'within a single \vebpage to estirnate its relevance to a query. But webpages are connected through h:yperlinks, and it is quite likely that there is a webpage containing the terrn 'search engine' that has a link to Yahoo's horne page. Can we use the inforrnation hidden in such links'?

Building on research in the sociology literature, an interesting analogy between links and bibliographic citations suggests a \vay to exploit link infoI'Ination: Just as influential authors and pubications are cited often, good webpages are likely to be often linked to. It is useful to distinguish between two types of pages, *authorities* and *hubs*. An authority is a page that is very relevant to a certain topic and that is recognized by other pages as authoritative on the subject. These other pages, called hubs, usually have a significant nUll1ber of hyperlinks to authorities, although they themselves are not very well known and do not necessarily carry a lot of content relevant to the given query. Hub pages could be cOlnpilatiol1s of resources about a topic on a site for professionals, lists of reco111mended sites for the hobbies of an individual user, or even a part of the bookIllarks of an individual user that are relevant to one of the user's interests; their Blain property is that they have IHany outgoing links to relevant pages. Good hub pages are often not well known and there may be few links pointing to a good hub. In contrast, good authorities are 'endorsed' by rnany good hubs and thus have many links froln good hub pages.

This symbiotic relationship between hubs and authorities is the basis for the HITS algoritlun, a link-based search algorithm that discovers high-quality pages that are relevant to a user's query terrns. The HITS algorithIll rnodels Web as a directed graph. Each webpage represents a node in the graph, and a hyperlink froIn page A to page B is represented as an edge between the two corresponding nodes.

Assulne that we are given a user query with several terIns. The algorithIll proceeds in two steps. In the first step, the *sarnpling step*, we collect a set of pages called the base set. The base set 11lOSt likely includes very relevant pages to the user's query, but the base set can still be quite large. In the second step, the *iteration step,* we find good authorities and good hubs arnong the pages in the base set.

The salnpling step retrieves a set of webpages that contain the query terrns, using sorne traditional technique. For exarnple, 'we can evaluate the query as a boolean keyword search and retrieve all webpages that contain the query terrns. We call the resulting set of pages the root set. The root set Inight not contain all relevant pages because senne authoritative pages rnight not include the user query \vords. But \ve expect that at least SOlne of the pages in the root set contain hyperlinks to the rnost relevant authoritative pages or that SCHne authoritative pages link to pages in the root set. This rnotivates our notion of a link page. We call a page a link page if it has a hyperlink to sorne page in the root set or if a page in the root set has a hyperlink to it. In order not to Iniss potentially relevant pages, we auglnent the root set by all link pages and we call the resulting set of pages the base set. Thus, the base set includes all

root pages and all link pages; we refer to a webpage in the base set as a base page.

Our goal in the second step of the algorithrn is to find out which base pages are good hubs and good authorities and to return the best authorities and hubs as the answers to the query. To quantify the quality of a base page as a hub and as an authority, we associate with each base page in the base set a hub weight and an authority weight. The hub weight of the page indicates the quality of the page as a hub, and the authority weight of the page indicates the quality of the page as an authority. We cornpute the weights of each page according to the intuition that a page is a good authority if rnany good hubs have hyperlinks to it, and that a page is a good hub if it has rnany outgoing hyperlinks to good authorities. Since we do not have any a priori knowledge about which pages are good hubs and authorities, we initialize all weights to one. We then update the authority and hub weights of base pages iteratively as described below.

Consider a base page p with hub weight h_p and with authority weight a_p. In one iteration, we update a_p to be the sum of the hub weights of all pages that have a hyperlink to p. Formally:

$$a_p = \sum_{\text{All base pages } q \text{ that have a link to } p} h_q$$

Analogously, we update h_p to be the sum of the weights of all pages that p points to:

$$h_p = \sum_{\text{All base pages } q \text{ such that } p \text{ has a link to } q} a_q$$

Cornparing the algorithrn with the other approaches to querying text that we discussed in this chapter, we note that the iteration step of the HITS algorithm—the distribution of the weights·· does not take into a,ccount the words on the base pages. In the iteration step, we are only concerned about the relationship between the base pages as represented by hyperlinks.

The lIlTS algorithrIl usually produces very good results. For exarnple, the five highest ranked results frorn Google ("rhich uses a variant of the HITS algorithrn) far the query 'Raghu Ramakrishnan' are the following webpages:

www.cs.wisc.edu/~raghu/raghu.html
www.cs.wisc.edu/~dbbook/dbbook.html
www.informatik.uni-trier.de/
~ley/db/indices/a-tree/r/Ramakrishnan:Raghu.html
www.informatik.uni-trier.de/

Computing bub and authority weights: We can use matrix notation to write the updates for all hub and authority weights in one step. Assume that we nUluber aU pages in the base set $\{1, 2, ..., n\}$. The adjacency matrix B of the base set is an n x n matrix whose entries are either Oor 1. The rnatrix entry $(i, \mathbf{j})$ is set to 1 if page i has a hyperlink to page **j**; it is set to 0 otherwise. We can also write the hub weights h and authority weights a in vector notation: h == $(\mathrm{h}_1, \ldots, h_n\rangle$ and $a = (\mathrm{al}, \text{'''} \ , a_n\rangle$. We can now rewrite our upda,te rules as follo\vs:

$$h = B \ . a, \quad \text{and} \quad a = B^T \ . h \ .$$

Unfolding this equation once, corresponding to the first iteration, we obtain:

$$h == BB^T h = (BBT)h, \quad \text{and} \quad a == BTBa = (BTB)a \ .$$

After the second iteration, we arrive at:

$$h = (BB^T)^2 h, \quad \text{and} \quad a = (B^T B)^2 a \ .$$

Results from linear algebra tell us that the sequence of iterations for the hub (resp. authority) weights converges to the principal eigenvectors of BET (resp. $B^T B$) if we normalize the weights before each iteration so that the suru of the squares of all weights is always 2 . n. Furthermore, results from linear algebra tell us that this convergence is independent of the choice of initial weights, as long as the initial weights are positive. Thus, our rather arbitrary choice of initial weights···--we initialized all hub and authority weights to 1—does not change the outcolne of the algorithm.

Google's Pigeon Rank: Google corIlputes the *pigeon rank (PRJ* for a webpage A using the following forrIlula, which is very sirnilar to the H.ub-Authority ranking functions:

$$PR(A) = (1 - d) + d(PR(T_1)/C(T_1) + ... + PR(T_n)/C(T_n))$$

'$\mathrm{T}_1 \ldots Tn$ are the pages that link (or 'point') to A, $C(T_i)$ is the rllllnber of links going out of page T_i, and d is a heuristically chosen constant (Google uses 0.85). Pigeon ranks fOflll a probability distribution over all webpages; the Slun of ranks over all pages is 1. If we consider a rnodel of user behavior in which a user randornly chooses a page and then repeatedly clicks on links until he gets bored and randoll1ly chooses a new page, the probability that the user visits a page is its Pigeon rank. The pages in the result of a search are ranked using a cornbination of an IR-style relevance l11etric and Pigeon rank.

SQL/MM: Full Text 'Full text' is described as data that can be searched, unlike simple character strings, and a new data type called FullText is introduced to support it. The Inethods associated 'with this type support searching for individual words, phrases, words that 'sound like' a query terlll, etc. Three 11lethods are of particular interest. CONTAINS checks if a FullText object contains a specified search terln (word or phrase). RANK returns the relevance rank of a FullText object with respect to a specified search terln. (I-Iow the rank is defined is left to the hnplementation.) IS ABOUT detennines whether the FullText object is sufficiently related to the specified search term. (The behavior of IS ABOUT is also left to the iInplelllentation.)

Relational DBMSs from IBIvI, Microsoft, and Oracle all support text fields, although they do not currently conforrll to the SQL/MM standard.

-ley/db/indices/a-tree/s/Seshadri:Praveen.html
www.acm.org/awards/fellows_citations_n-z/ramakrishnan.htmI

The first result is Rarnakrishnan's horne page; the second is the horne page for this book; the third is the page listing his publications in the popular DBLP bibliography; and the fourth (initially puzzling) result is the list of publications for a forrner student of his.

27.5 MANAGING TEXT IN A DBMS

In preceding sections, we saw how large text collections are indexed and queried in JR, systerns and Web search engines. We now consider the additional challenges raised by integrating text data into database systerns.

The basic approach being pursued by the SQI.I standards cornrnunity is to treat text docllrnents as a new data type, FullText, that can appear as the value of a field in a table. If we define a table with a single cohunn of type FullText, each row in the table corresponds to a docurnent in a dOCllllnent collection. IVlethods of FullText can be llsed in the WHERE clause of SQL queries to retrieve rows containing text objects that Inatch an IR-style search criterion. The relevance rank of a FullText object can be explicitly retrieved using the RANK rnethod, and this can be llsed to sort results by relevance.

Several points ruust be kept in rnind as we consider this approach:

- This is an extremely general approach, a,nel the perforlnance of a SQL systern that supports such an extension is likely to be inferior to a specialized IR SystCIll.

- The rnodel of data does not ad.equately reflect docurnents with additional rnetadata. If we store docurnents in a table with a FullText colurnn and use additional cohl1nns to store rnetadata--for exarnple, author, title, SUlll-Inary, rating, popularity—relevance rneasures that cornbine nletadata 'with IR similarity measures 11lUSt be expressed using lle\V user-defined rneth-ods, because the RANK rnethod only has access to the FullText object, and not the rnetadata. The ernergence of XML docurnents, which have non-uniforrn, partial rlletadata, further cornplicates nlatters.
- The handling of updates is unclear. As we have seen, IR indexes are corll-plex, and expensive to 111aintain. Requiring a systern to update the indexes before the updating transaction cOl1ullits can irnpose a severe perfonnance penalty.

27.5.1 Loosely Coupled Inverted Index

The irrlplenlcntation approach used in current relational DBMSs that support text fields is to have a separate text-search engine that is loosely coupled to the DBMS. The engine periodically updates the indexes, but provides no transac-tional guarantees. Thus, a transaction could insert (a row containing) a text object and cornrnit, and a subsequent transaction that issues a Inatching search might not retrieve the (row containing the) object.

27.6 A DATA MODEL FOR XML

.Aswe saw in Section 7.4.1, XML provides a way to rnark up a docurnent with rneaningful tags that irnpart SaIne partial structure to the docurnent. *Semistructured data rnodels,* which we introduce in this section, capture rnuch of the structure in XML doculnents, while abstracting away Inany deta.ils.[1] Sernistructured data Inodels have the potential to serve as a forInal foundation for XlVIL and enable us to rigorously define the sernantics of queries over XlVIL, which we discuss in Section 27.7.

27.6.1 Motivation for Loose Structure

Consider a set of doculnents on the Web that contain hyperlinks to other doc-UHlents. These docurnents, although not eornpletely unstructured, cannot be rnodeled naturally in the relational data rnodel because the pattern of hyper-links is not regular across docurnents. In fact, every HTML docurnent has

1.An iruportant aspect of XML tha.t is *not* captured is the ordering of elements. A more complete data model called XData has been proposed by the W3C committee that is developing XML standards, but we do not discuss it here.

XML Data Models: 'A number of data lnodels for XML are being considered by standards COilllnittees such as ISO and W3C. W3C's Infoset is a tree-structured model, and each node can be retrieved through an accessor function. A version called Post-Validation Infoset (PSVI) serves as the data model for XML Schelna. TheXQuery language has yet another data model associated with it. The plethora of l110dels is due to parallel developrnent in SOllle cases, and due to different objectives in others. Nonetheless, all these nlodels have loosely-structured trees as their central feature.

some minirnal structure, such as the text in the TITLE tag versus the text in the docunlent body, or text that is highlighted versus text that is not. As another example, a bibliography file also has a certain degree of structure due to fields such as *author* and *title,* but is otherwise unstructured text. Even data that is 'unstructured', such as free text or an ilnage or a video clip, typically has some associated information such as timestamp or author infornlation that contributes partial structure.

We refer to data with such partial structure as semistructured data. There are rnany reasons why data might be semistructured. First, the structure of data nlight be irnplicit, hidden, unknown, or the user Inight choose to ignore it. Second, when integrating data froln several heterogeneous sources, data exchange and transforrnation are inlportant problerns. We need a highly flexible data rnodel to integrate data froIn all types of data sources including flat files and legacy systenls; a structured data model such as the relational rnodel is often too rigid. Third, we cannot query a structured database without knowing the scheIna, but sOlnetimes we want to query the data without full knowledge of the scherna. For exarnple, we cannot express the query "Where in the database can we find the string *Malgudi*?" in a relational database systern \vithout knowing the schcrna, and knowing which fields contain such text values.

27.6.2 A Graph Model

All data rnodels proposed for sernistrnctured data represent the data as scnne kind of labeled graph. Nodes in the graph correspond to cornpound objects or atornic values.. Each edge indicates an object-subobject or object-value relationship. Leaf nodes, i.e, nodes with no outgoing edges have a value a.ssociatecl \vith thern. rrhere is no separate scherna and no auxiliary description; the data in the graph is self describing. For exarnple, consider the graph shown in Figure 27.7, which represents part of the XML data f1'01n Figure 7.2. The root node of the graph represents the outennost elernent, BOOKLIST. The node has three children that are labeled with the elClnent narne BOOK, since the list of books

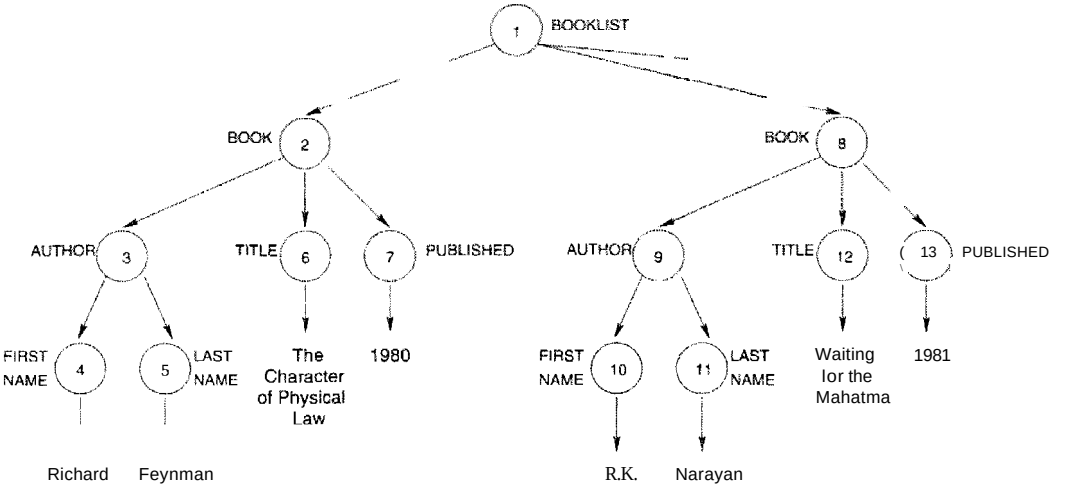


Figure 27.7 The Semistructured Data Model

consists of three individual books. The numbers within the nodes indicate the object identifier associated with the corresponding object.

We now describe one of the proposed data models for semistructured data, called the object exchange model (OEM). Each object is described by a quadruple consisting of a *label,* a *type,* the *value* of the object, and all. *object identifier* which is a unique identifier for the object. Since each object has a label that can be thought of as a column nallle in the relational model, and each object has a type that can be thought of as the column type in the relational rnodel, the object exchange Illodel is self-describing. Labels in OEM should be as infol'rnative as possible, since they serve two purposes they can be used to identify an object as well as to convey the meaning of an object. For example, we can represent the last HaIne of an author as follows:

(lastName, string, "Feynman")

More cOInplex objects are decornposed hierarchically into srnaller objects. For exalllple, a,n author naIne can contain a first narne and a last narne. rrhis object is described as follows:

(authorName, set, *{fiTstname*$_1$, *lastnaTnel})*
firstname$_1$ is (firstName, string, "Richard")
lastname$_1$ is (lastName, string, "Feynman")

As another exarnple, an object representing a set of books is described as follows:

(bookList, set, $\{book_1, book_2, book_3\}\rangle$
book} is (book, set, $\{author_1, title\}, published_1\}\rangle$

SQL and XML: XQuery is a standard proposed by the World-Wide Web Consortiurn (W3C). In parallel, standards con:unittees developing the SQL standards have been working on a successor to SQL:1999 that supports XML. The part that relates to XML is tentatively called SQL/XML and details can be found at http://sqlx.arg.

book$_2$ is ⟨book, set, {*author*$_2$, *title*$_2$, *published*$_2$}⟩
book$_3$ is ⟨book, set, {*author*$_3$, *t'itle3*, *Published3*})
authoT3 is (author, set, {*!'lrstnarnJe3*, *lastname*$_3$}⟩
t'itle3 is (title, string, liThe English Teacher")
published$_3$ is (published, integer, 1980)

27.7 XQU'ERY: QUERYING XML DATA

Given that XlvIL.. doculnents are encoded in a way that reflects (a considerable amount of) structure, we have the opportunity to use a high-level language that exploits this structure to conveniently retrieve data fro111 within such documents. Such a language would also allow us to easily translate XML data between different DTDs, as we lllUSt when integrating data from multiple sources. At the tirne of writing of this book, XQuery is the W3C standard query language for XML data. In this section, we give a brief overview of XQuery.

27.7.1 Path Expressions

Consider the XlvIL.. dOCUlnent shown in Figure 7.2. The following exarnple query returns the last nanles of all authors, assullling that our XML docurnent resides at the location www.ourbookstore.com/books.xml.

```
FOR
   $1 IN doc(www.ourbookstore.com/books.xml)//AUTHOR/LASTNAME
RETURN <RESULT> $1 </RESULT>
```

This exarnple illustrates sonle of the basic constructs of XQuery. The FOR clause in XQuery is roughly analogous to the FROM clause in SQL. The RETURN clause is sirnilar to the SELECT clause. We return to the general fornl of queries shortly, after introducing an irnportant concept called a path expression.

The expression

```
doc(www.ourbookstore.com/books.xml)//AUTHOR/LASTNAME
```

XPath and Other XML Query Languages: Path expressions in XQuery are derived frorn XPath, an earlier XML query facility. Path expressions in XPath can be qualified with selection conditions, and can utilize several built-in functions (e.g., counting the nurnber of nodes rnatched by the expression). Many of XQuery's features are borrowed {roIn earlier languages, including XML-QL and Quilt.

in the FOR clause is an exarnple of a path expression. It specifies a path involving three entities: the docurnent itself, the AUTHOR elernents and the LASTNAME elernents.

The path relationship is expressed through separators / and //. The separator // specifies that the AUTHOR elernent can be nested anywhere within the document whereas the separator / constrains the LASTNAME elernent to be nested immediately under (in terms of the graph structure of the docurnent) the AUTHOR element. Evaluating a path expression returns a *set* of elernents that rnatch the expression. The variable *l* in the example query is bound in turn to each LASTNAME elernent returned by evaluating the path expression. (To distinguish variable naInes fronl normal text, variable narnes in XQuery are prefixed with a dollar sign $.)

The RETURN clause constructs the query result-·--which is also an XML docurnent-···_· by bracketing each value to which the variable *l* is bound with the tag RESULT. If the exanlple query is applied to the sarnple data shown in Figure 7.2, the result would be the following XML docurnent:

```
<RESULT><LASTNAME>Feynman </LASTNAME></RESULT>
<RESULT><LASTNAME>Narayan </LASTNAME></RESULT>
```

We use the docurnent in Figure 7.2 as our input in the rest of this chapter.

27.7.2 FLWR Expressions

The basic fornl of an XQuery consists of a FLWR **expression**, where the letters denote the **FOR**, **LET**, WHERE and RETURN clauses. The FOR and LET clauses bind variables to values through path expressions. These values are qualified by the WHERE clause, and the result XML fragrnent is constructed by the RETURN clause.

rrhe difference between a FOR and LET clause is that while FOR binds a variable to each elernent specified by the path expression, LET binds a variable to the whole *collection* of elernents. Thus, if we change our exarnple query to:

```
LET
   $IINdoc(www.ourbookstore.com/books.xm1)//AUTHOR/LASTNAME
RETURN <RESULT> $1 </RESULT>
```

then the result of the query beconles:

```
<RESULT>
   <LASTNAME>Feynman</LASTNAME>
   <LASTNAME>Narayan</LASTNAME>
</RESULT>
```

Selection conditions are expressed using the WHERE clause. Also, the output of a query is not lirnited to a single elernent. These points are illustrated by the following query, which finds the first and last names of all authors who wrote a book that was published in 1980:

```
FOR $b IN doc(www.ourbookstore.com/books.xm1)/BOOKLIST/BOOK
WHERE $b/PUBLISHED='19S0'
RETURN
   <RESULT> $b/AUTHOR/FIRSTNAME, $b/AUTHOR/LASTNAME </RESULT>
```

The result of the above query is the following XML docurnent:

```
<RESULT>
   <FIRSTNAME>Richard </FIRSTNAME><LASTNAME>Feynman </LASTNAME>
</RESULT>
<RESULT>
   <FIRSTNAME>R.K. </FIRSTNAME><LASTNAME>Narayan </LASTNAME>
</RESULT>
```

For the specific DTI) in this exalnple, where a BOOK elernent has only one AUTHOR, the above query can be written by using a different path expression in the FOR clause, as follows.

```
FOR $a IN
   doc(www.ourbookstore.com/books.xml)
      /BOOKLIST/BOOK[PUBLISHED='19S0']/AUTHOR
RETURN <RESULT> $a/FIRSTNAME, $a/LASTNAME </RESULT>
```

rrhe path expression in this query is an instance of a branching path expression. The variable / is now bound to every AUTHOR elernent that rnatches the path doc/BOOKLIST/BOOK/AUTHOR where the intennediate BOOK elClnent is constrained to have a PUBLISHED elernent nested inunediately within it with the value 1980.

27.7.3 Ordering of Elements

XML data consists of *ordered* doculnents and so the query language IllUSt return data in SOUle order. The selnantics of XQuery is that a path expression returns results sorted in document order. Thus, variables in the FOR clause are bound in doculnent order. If however, we desire a different order, we can explicitly order the output as shown in the follo\ving query, which returns TITLE elernents sorted lexicographically.

```
FOR
   $b IN doc(www.ourbookstore.com/books.xml)/BOOKLIST/BOOK
RETURN <BOOKTITLES> $b/TITLE </BOOKTITLES>
SORT BY TITLE
```

27.7.4 Grouping and Generation of Collection Values

Our next example illustrates grouping in XQuery, which allows us to generate a new collection value for each group. (Contrast this with grouping in SQL, which only allows us to generate an aggregate value (e.g., SUM) per group.) Suppose that for each year we want to find the last narnes of authors who wrote a book published in that year. We group by year of publication and generate a list of last names for each year:

```
FOR $p IN DISTINCT
   doc(www.ourbookstore.com/books.xml)/BOOKLIST/BOOK/PUBLISHED
   RETURN
   <RESULT>
      $p,
      FOR $a IN DISTINCT /BOOKLIST/BOOK[PUBLISHED=$pJ/AUTHOR
         RETURN $a
   </RESULT>
```

The keyword *DISTINCT* elirninates duplicates fronl the collection returned by a, path expression. Using the XML docurnent in Figure 7.2 as input, the above query produces the following result:

```
<RESULT> <PUBLISHED>1980</PUBLISHED>
   <LASTNAME>Feynman</LASTNAME>
   <LASTNAME>Narayan</LASTNAME>
</RESULT>
<RESULT> <PUBLISHED>1981</PUBLISHED>
   <LASTNAME>Narayan</LASTNAME>
</RESULT>
```

27.8 EFFICIENT EVALUATION OF XML QUERIES

X.Query operates on XML data and produces XTvIL data as output. In order to be able to evaluate queries efficiently, we need to address the follo\ving issues.

- Storage: We can use an existing storage systerIl like a relational or object oriented systerll or design a new storage forInat for XML doclunents. There are several ways to use a relational systenl to store XML. One of thern is to store the XML data as Character Large Objects (CLOBs). (CLOBS were discussed in Chapter 23.) In this case, however, we cannot exploit the query processing infrastructure provided by the relational systerrl and would instead have to process XQuery outside the database systenl. In order to circumvent this problenl, we need to identify a scherna according to which the XML data can be stored. These points are discussed in Section 27.8.1.

- Indexing: Path expressions add a lot of richness to XQuery and yield lllany new access patterns over the data. If we use a relational system for storing XML data, then we are constrained to use only relational indexes like the B-Tree. However, if we use a native storage engine, then we have the option of building novel index structures for path expressions, some of which are discussed in Section 27.8.2.

- Query Optimization: Optirnization of queries in XQuery is an open problern. The work so far in this area can be divided into three parts. 'rhe first is developing an algebra for XQuery, analogous to relational algebra. The second research direction is providing statistics for path expression queries. Finally, SOlne work has addressed sirnplification of queries by exploiting constraints on the data. Since query optirnization for X.Query is still at a prelirninary stage, we do not cover it in this chapter.

Another issue to be considered while designing a new storage systeul for XML data is the verbosity of repeated tags. As we see in Section 27.8.1) using a relational storage systelu addresses this problern since tag narnes are not stored repeatedly. If on the other hand, we want to build a native storage systcrn, then the rnanner in which the XML data is cornpressed becornes significant. Several cornpression. algorithrHs are known that achieve cOlnpression ratios close to relational storage, 1n1t we do not discuss ther11 here.

27.8.1 Storing XML in RDBMS

One natural candidate for storing XML data is a relational database systern. The ruain issues involved in storing XML data in a relational systelTI are:

Commercial database systems and XML: Many relational and object-relational database systerll vendors are currently looking into support for XML in their database engines. Several vendors of object-oriented database Inanagenlent systems already offer database engines that can store XML data whose contents can be accessed through graphical 11ser interfaces or Jl server-side Java extensions.

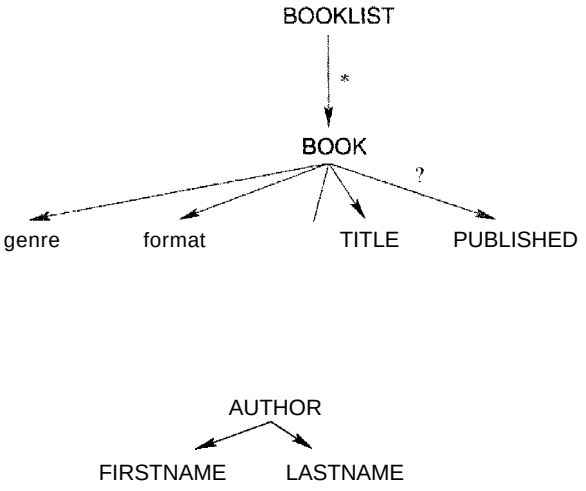


Figure 27.8 Bookstore XML DTD Element Relationships

- *C;hoice of relational schema*: In order to use an RDBMS, we need a scherna. What relational schema should we use even assuming that the XML data COUles with an associated scherna?
- *Queries:* Queries on XML data are in XQuery whereas a relational systern can only handle SQL. Queries in XQuery therefore need to be *translated* into SQL.
- *Reconstruction*: rrhe output of XQuery is XML. Thus, the result of a SQL query needs to be converted back into XML.

Mapping XML Data to Relations

We illustrate the rnapping process through our bookstore exarnple. rrhe nesting rela,tionships among the different elernents in the DTD is shown in Figure 27.8. The edges indicate the nature of the nesting.

()ne way to derive a l'clation.al schelna is as follows. We begin at the BOOKLIST elernent and create a relation to store it. rrraversing down froIn BOOKLIST, we get BOOK following a * edge. This edge indicates that we store the BOOK elernents in a separate relation. Traversing further down, we see that all elcrnents and

attributes nested within BOOK occur at 1l10St once. Hence, we can store thenl in the saIne relation as BOOK. The resulting relational schelIla *Relschernal* is shown below.

BOOKLIST(*id:* integer)
BOOK *(booklistid:* integer, *author_firstname*: string,
author_lastnarne: string, *title:* string,
published: string, *genre:* string, *format*: string)

BOOK. *booklistid* connects BOOK to BOoKLIST. Since a DTD has only one base type, string, the only base type used in the above schelna is string. The constraints expressed through the DTD are expressed in the relational schema. For instance, since every BOOK must have a TITLE child, we Illust constrain the *title* column to be *non-null*.

Alternatively, if the DrrD is changed to allow BOOK to have more than one AUTHOR child, then the AUTHOR elements cannot be stored in the sallie relation as BOOK. This change yields the following relational schema *Relschema2*.

BOOKLIST(*id:* integer)
BOOK *(id:* integer, *booklistid:* integer,
title: string, *published:* string, *genre:* string, *for1nat:* string)
AUTHoR(*bookid:* integer, *firstname:* string, *lastname:* string)

The column AUTHOR. *bookid* connects AUTHOR to BOOK.

Query Processing

Consider the following example query again:

```
FOR
   $b IN doc(www.ourbookstore.com/books.xml)/BOOKLIST/BooK
WHERE $b/PUBLISHED='1980'
RETURN
   <RESULT> $b/AUTHOR/FIRSTNAME, $b/AUTHOR/LASTNAME </RESULT>
```

If the nlapping between the XML data and relational tables is known, then we can construct a SQL query that returns all colunlIls that are needed to reconstruct the result XML docuIIlent for this query. Conditions enforced by the path expressions and the WHERE clause are translated into equivalent conditions in the SQL query. We obtain the following equivalent SQL query if we use *Relschema1* as our relational scherna.

```
SELECT BOOK.author_firstname, BOOK.author_lastname
```

```
FROM    BOOK, BOOKLIST
WHERE   BOOKLIST.id = BOOK.booklistid
          AND BOOK.published='1980'
```

The results thus returned by the relational query processor are then tagged, outside the relational systern, as specified by the RETURN clause. This is the result of the *reconstruct'ion* phase.

In order to understand this better, consider what happens if we allow a BOOK to have 111ultiple AUTHOR children. Assume that we use *Rel8chema2* as our relational schema. Processing the FOR and WHERE clauses tells us that it is necessary to join relations BOOKLIST and BOOK with a selection on the BOOK relation corresponding to the year condition in the above query. Since the RETURN clause needs information about AUTHOR elements, we need to further join the BOOK relation with the AUTHOR relation and project the *jir8tname* and *lastname* columns in the latter. Finally, since each binding of the variable $b in the above query produces one RESULT element, and since each BOOK is now allowed to have more than one AUTHOR, we need to project the *id* column of the BOOK relation. Based on these observations, we obtain the following equivalent SQL query:

```
SELECT    BOOK.id, AUTHOR. firstname , AUTHOR.lastname
FROM      BOOK, BOOKLIST, AUTHOR
WHERE     BOOKLIST.id = BOOK.booklistid AND
            BOOK.id = AUTHOR.bookid AND BOOK.published='1980'
GROUP BY BOOK.id
```

T'he result is grouped by BOOK.id. The tagger outside the database system now receives results clustered by the BOOK element and can tag the resulting tuples on the fly.

Publishing Relational Data as XML

Since XML has elnerged as the standard data exchange forrnat for business applications, it is necessary to publish existing business data as XML. Most operational business data is stored in relational systerns. Consequently, 111echanisrns have been proposed to publish such data as XML docul11ents. These involve a language for specifying henv to tag and structure relational data and an irnplernentation to carry out the conversion. This 111apping is in some sense the reverse of the XML-to-relationaJ rnapping used to store XML data. 'The conversion process Inirnics the reconstruction phase when we execute XQuery using a relational systern. The published XML data can be thought of as an XML view of relational data. This view can be queried using XQuery. One

lnethod of executing XQuery on such vie'ws is to translate thCIll into SQL and thCIl construct the XML result.

27.8.2 Indexing XML Repositories

Path expressions are at the heart of all proposed XIVIL query languages, in particular XQuery. A natural question that arises is how to index XML data to support path expression evaluation. The ainl of this section is to give a flavor of the indexing techniques proposed for this probleul. We consider the OEM rnodel of senlistructured data, 'where the data is self-describing and there is no separate scherna.

Using a B+ Tree to Index Values

Consider the following XQuery exaluple, which we discussed earlier on the bookstore XML data in Figure 7.2. The OEM representation of this data is shown in Figure 27.7.

```
FOR
   $b IN doc(www.ourbookstore.com/books.xml)/BOOKLIST/BOOK
WHERE  $b/PUBLISHED='1980'
RETURN
   <RESULT>  $b/AUTHOR/FIRSTNAME,  $b/AUTHOR/LASTNAME  </RESULT>
```

This query specifies joins allIong the objects with labels BOOKLIST, BOOK, AUTHOR, FIRSTNAME, LASTNAME and PUBLISHED with a selection condition on PUBLISHED objects.

Let us suppose that we are evaluating this query in the absence of any indexes for path expressions. However, we do have a value index such as a B-T'ree that enables us to find the ids of all objects with label PUBLISHED and value 1980. There are several ways of executing this query under these assumptions.

For instance, we could begin at the docurncnt root and traverse down the data graph through the BOOKLIST object to the BOOK objects. By further traversing the data graph downwards, for each BOOK object we can check whether it satisfies the valuc'predicate (PUBLISHED=‘1980’). Finally, for those BOOK objects that satisfy the predicate, we can find the relevant FIRSTNAME and LASTNAME objects. This approach corresponds to a top-down evaluation of the query.

Alternatively, we could begin by using the value index to find all PUBLISHED ol)jects that satisfy PUBLISHED='1980'. If the data graph can be traversed in the reverse directiono·that is, given an object, we can find its parent—then we

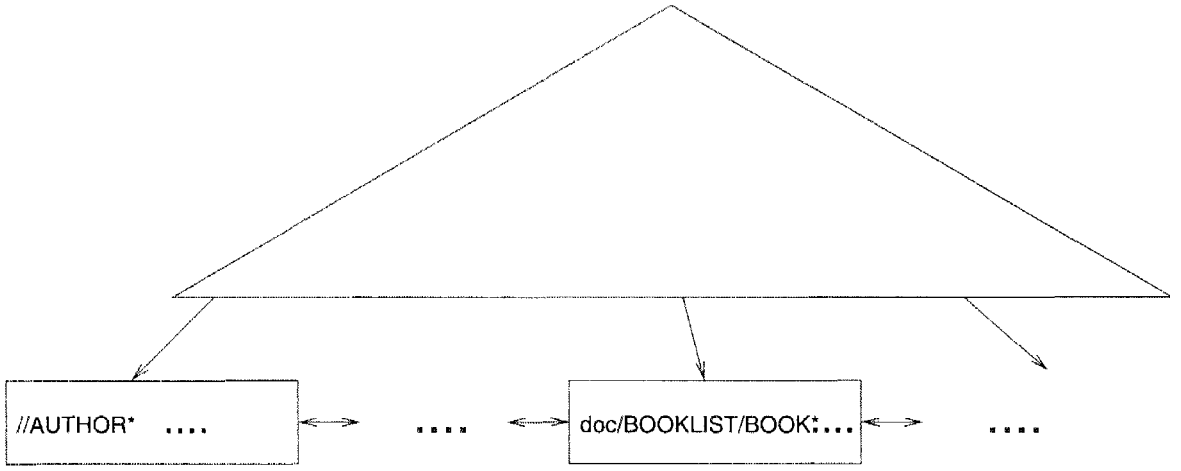


Figure 27.9 Path Expressions in a B-Tree

can find all parents of the PUBLISHED objects retaining only those that have label BOOK. We can continue in this manner until we find the FIRSTNAME and LASTNAME objects of interest. Observe that we need to perforrll all joins in the query on the fly.

Indexing on Structure vs. Value

Now let us ask ourselves whether traditional indexing solutions like the B-Tree can be used to index path expressions. We can use the B-Tree to rllap a path expression to the ids of all objects returned by it. The idea is to treat all path expressions as strings and order therIl lexicographically. Every leaf entry in the B-Tree contains a string representing a, path expression and a list of ids corresponding to its result. Figure 27.9 shows how such a B-Tree would look. Let us contrast this with the traditional problern of indexing a well-ordered dornain like integers for point queries. In the latter case, the nurnber of distinct point queries that can be posed is just the rnllnber of data values and so is linear in the data size.

The scenario with path indexing is fundarnentally different—the variety of ways in which we can cornbine tags to forrn (sirnple) path expressions C011-pled with the power of placing // separators leads to a rnuch larger nurnber of possible path expressions. For instance, an AUTHOR clcrnent in the exarn-pIe in Figure 27.7 is returned as part of the qllcries BOOKLIST/BOOK/AUTHOR, / / AUTHOR, / /BOOK/ / AUTHOR, BOOKLIST/ / AUTHOR and so OII. The nurnber of distinct queries can in fact be exponential in the data size (lneasured in tenns of the rnunber of XIVIL elelnents) in the worst case. This is \vhat rnotivates the search for alternative strategies to index path expressions.

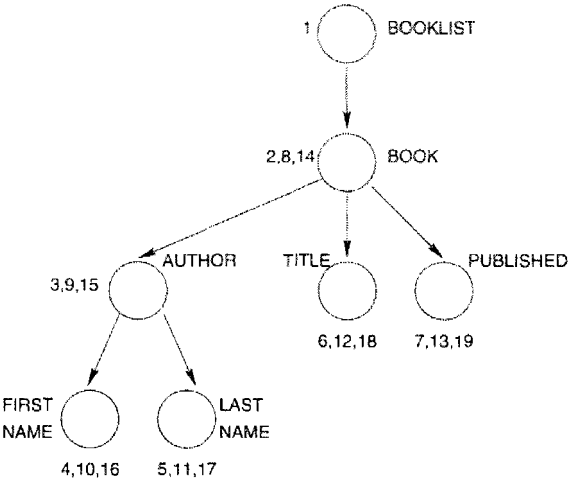


Figure 27.10 Example Path Index

The approach taken is to represent the mapping between a path expression and its result by means of a structural sunlIllary which takes the fornl of another labeled, directed graph. rrhe idea is to preserve all the paths in the data graph in the surllrnary graph, while having far fewer nodes and edges. An extent is associated with each node in the SUllllnary. The extent of an index node is a subset of the data nodes. The surnmary graph along with the extents constitutes a path index. A path expression is evaluated using the index by evaluating it against the sumrnary graph and then taking the union of the extents of all rnatching nodes. This yields the index result of the path expression query. The index covers a path expression if the index result is the eorrect result; obviously, we can use an index to evaluate a path expression only if the index covers it.

Consider the structural SUlnrnary shown in Figure 27.10. rrhis is a path index for the data in Figure 27.7. Tlhe nurnbers shown beside the nodes correspond to the respective extents. Let us now exarnine how this index can change the top-down evaluation of the exaruple query used earlier to illustrate B+ tree value indexes.

rrhe top-down evaluation as outlined above begins at the docurnent root and traverses down to the BOOK objects. rrhis can be achieved rnore efficiently by the path index. Instead of traversing the data graph, we can traverse the path index down to the BOOK object in the index and look up its extent, which gives us the ids of all BOOK objects that rnatch the path expression in the FOR clause. The rest of the evaluation then proceeds as before. Thus, the path index saves us frorn perfonning joins by essentially precorIlputing thern. We note here that the path index shown in Figure 27.10 is isornorphic to the DTD schcIIla graph shO\vII in Figure 27.8. This drives horne the point that the path index \vithout the extents is a structural SUHllnary of the data.

rrhe ahove path index is the Strong Dataguide. If we treat path expressions as strings, then the dataguide is the trie representing thern. The trie is a well-known data structure used to search regular expressions over text. This shows the deeper unity between the research on indexing text and the XML path indexing work. Several other path indexes have been also proposed for senli-structured data, and this is an active area of research.

27.9 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What is information retrieval? (Section 27.1)
- What are some of the differences between DBMS and IR systems? Describe the differences between a ranked query and a boolean query. (Section 27.2)
- What is the vector space model, and what are its advantages? (Section 27.2.1)
- What is TF/IDF terrn weighting, and why do we weigh by both? We do we eliminate stop words? What is length norrnalization, and why is it done? (Section 27.2.2)
- How can we measure document similarity? (Sections 27.2.3)
- What are *precision* and *recall,* and how do they relate to each other? (Section 27.2.4)
- Describe the following two index structures for text: Inverted index and signature file. What is a bit-sliced signature file? (Section 27.3)
- How are web search engines architected? Ilow does the "hubs and authorities" a.lgorithrn work? Can you illustrate it on a srnall set of pages? (Section 27.4)
- What support is there for rnanaging text in a DBMS? (Section 27.5)
- Descibe the OEM data rnodel for sernistructured data. (Section 27.6)
- What are the elernents of XQuery? What is a path expression? What is an FLWR expression? How can we order the output of query? flow do we group query outputs? (Section 27.7)
- Describe how XTvIL data can be stored in a relational DBMS. How do we rna,p XML data to relations? Can we use the query processing infrastructure of the relational DBIvIS? How do 'we publish relational data as XML? (Section 27.8.1)

- I-Iow do we index collections of XML doeunlents? What is the difference between indexing on structure versus indexing on value? What is a path index'? (Section 27.8.2)

EXERCISES

Exercise 27.1 Carry out the following tasks.

1. Given an ASCII file, cOInpute the frequency of each word and create a plot siInilar to Figure 27.3. (Feel free to use public dornain plotting software.) Run the progralll on the collection of files currently in your directory and see whether the distribution of frequencies is Zipfian. How can you use such plots to create lists of stop words?
2. The Porter stenliller is widely used, and code irnplernenting it is freely available. Download a copy, and run it on your collection of docuInents.
3. One criticisIn of the vector space nlodel and its use in sirnilarity checking is that it treats tenns as occurring independently of each other. In practice, Inany words tend to occur together (e.g., ambulance and emergency). Write a program that scans an ASCII file and lists all pairs of words that occur within 5 words of each other. For each pair of words, you now have a frequency, and should be able to create a plot like Figure 27.3 with pairs of words on the X-axis. Run this program on some sample docuIll€nt collections. What do the results suggest about co-occurrences of words?

Exercise 27.2 Assunle you are given a docurnent database that contains SIX documents. After stemming, the docurnents contain the following ternlS:

Document	Terrns
1	car rnanufacturer Honda auto
2	auto cornputer navigation
3	Honda navigation
4	11lanufactllrer cOlnputer IBM
5	IBNI personal cOInputer
6	car Beetle VW

Answer the following questions.

1. 8ho\v the result of creating an inverted file on the docurncnts.
2. Show the result of creating a signature file with a width of 5 bits. Construct your own hashing function that rnaps terms to bit positions.
3. Evaluate the following boolea.n queries using the inverted file and the signature file that you created: 'car', 'IBM' AND 'COIuputer', 'IBM' AND 'car', 'IBM' OR 'auto', and 'IBM' AND 'computer' AND 'rnanufacturer'.
4. Assurne that the query loacl against the docurnent databa.se consists of exactly the queries that were stated in the previous question. Also assume that each of these queries is evaluated exactly onCC.

   (a) Design a signature file with a width of 3 bits and design a hashing function that minimizes the overall nurnber of false positives retrieved when evaluating the

(b) Design a signature file with a width of 6 bits and a hashing function that minimizes the overall nUlnber of false positives.

(c) Assume you want to construct a signature file. What is the sInallest signature width that allows you to evaluate all queries without retrieving any false positives?

5. Consider the following ranked queries: 'car, 'IBM COIllputer', 'IBM car', 'IBM auto', and 'IBM COIllputer rnanufacturer'.

(a) Calculate the IDF for every tenn in the database.

(b) For each doculnent, show its doctunent vector.

(c) For each query, calculate the relevance of each doclunent in the database, with and without the length norrnalization step.

(d) Describe how you would use the inverted index to identify the top two documents that Illatch each query.

(e) How would having the inverted lists sorted by relevance instead of document id affect your answer to the previous question?

(f) Replace each docurnent with a variation that contains 10 copies of the same document. For each query, recompute the relevance of each document, with and without the length normalization step.

Exercise 27.3 Assume you are given the following steIIllned docurnent database:

Document	Terms
1	car car IIlanufacturer car car Honda auto
2	auto computer navigation
3	Honda navigation auto
4	manufacturer computer IBl\II graphics
5	IBM personal IBM computer IBl\II IBl\I! IBM IBM
6	car Beetle VW Honda

Using this database, repeat the previous exercise.

Exercise 27.4 You are in charge of the Genghis ('We execute fast') search engine. You are designing your server cluster to handle 500 Inillion hits a day and 10 billion pages of indexed data. Each rnachine costs $1000, and can store 10 million pages and respond to 200 queries per second (against these pages).

1. If you were given a budget of $500,000 dollars for purchasing Inachines, and were required to index all 10 billion pages, could you do it?
2. What is the IllinirIlurIl budget to index all pages? If you assurne that each query can be answered by looking at data in just one (10 rnillion page) partition, and that queries are uniformly distributed across partitions, what peak load (in nuruber of queries per second) can such a cluster handle?
3. How would your answer to the previous question change if each query, on average, accessed two partitions?
4. What is the ruinirlllnl1 budget required to handle the desired load of 500 rnillion hits per day if all queries are on a *single partition*? Assurne that queries are uniforrnly distributed with respect to tirTle of day.

5. How would your answer to the previous question change if the rllllnher of queries per day went up to 5 billion hits per day? How would it change if the number of pages went up to 100 billion'?
6. Assurne that each query accesses just one partition, that queries are ullifonnly distributed across partitions, but that at any given time, the peak load on a partition is upto 10 times the average load. What is the rminiIlHlnl budget for purchasing Inachines in this scenario?
7. Take the cost for rnachines [raIn the previous question and rnultiply it by 10 to reflect the costs of Illaintenance, adrninistration, network bandwidth, etc. This anlount is your annual cost of operation. Assume that you charge advertisers 2 cents per page. What fraction of your inventory (i.e., the total nUlllber of pages that you serve over the course of a year) do you have to sell in order to make a profit?

Exercise 27.5 Assume that the base set of the HITS algorithrn consists of the set of Web pages displayed in the following table. An entry should be interpreted as follows: Web page 1 has hyperlinks to pages 5 and 6.

Webpage	*Pages that this page has links to*
1	5, 6, 7
2	5, 7
3	6, 8
4	
5	1, 2
6	1, 3
7	1, 2
8	4

1. Run five iterations of the HITS algorithlll and find the highest ranked authority and the highest ranked hub.
2. Cornpute Google's Pigeon Rank for each page.

Exercise 27.6 Consider the following description of itelllS shown in the Eggface cornputer rnail-order catalog.

"Eggface sells hardware and software. We sell the new PalIn Pilot V for $400; its part nUlnber is 345. We also sell the IBM ThinkPad 570 for only $1999; its part nUIllber is 3784. We sell both business and entertainrnent software. I:vlicrosoft Office 2000 has just arrived and you can purchase the Standard Edition for only $140, part number 974; the Professional Edition is $200, part 975. '1'he new desktop publishing software from Adobe called InDesign is here for only $200, part 664. We carry the newest gaInes from Blizzard software. You can start playing Diablo II for only $30, part nurnber 12, and yon can purchase Starcraft for only $10, part nlllIlber 812. Our goal is cornplete cllstorner satisfaction·····-ifwe don't have what you want in stock, we'll give you SIO off your next purchase!"

1. Design an HTML docllrnent that depicts the itelllS offered by Eggface.
2. Create a well-formed XML doculnent that describes the contents of the Eggface catalog.
3. Create a TYr.D for your XML docurnent and rnake sure that the docuJnent you created in the last question is valid with respect to this *1Y1'1)*,

4. Write an XQuery query that lists all software items in the catalog, sorted by price.
5. Write an XQuery query that, for each vendor, lists all software iterl1s froIn that vendor (i.e., one row in the result per vendor).
6. Write an XQuery query that lists the prices of all hardware items in the catalog.
7. Depict the catalog data in the semistructured data model as shown in Figure 27.7.
8. Build a dataguide for this data. Discuss how it can be used (or not) for each of the above queries.
9. Design a relational schellla to publish this data.

Exercise **27.7** A university database contains infonnation about professors and the courses they teach. The university has decided to publish this information on the Web and you are in charge of the execution. You are given the following information about the contents of the database:

In the fall sernester 1999, the course 'Introduction to Database Management Systems' was taught by Professor Ioannidis. The course took place Mondays and Wednesdays from 9–10 a.m. in room 101. The discussion section was held on Fridays fTOIn 9-10 a.m. Also in the fall semester 1999, the course 'Advanced Database Management Systems' was taught by Professor Carey. Thirty five students took that course which was held in room 110 Tuesdays and Thursdays from 1–2 p.m. In the spring semester 1999, the course 'Introduction to Database Management Systems' was taught by U.N. Owen on Tuesdays and Thursdays frOIn 3-4 p.m. in room 110. Sixty three students were enrolled; the discussion section was on Thursdays from 4–5 p.m. The other course taught in the spring semester was 'Advanced Database Management Systems' by Professor Ioannidis, Monday, Wednesday, and Friday frorn 8-9 a.m.

1. Create a well-formed XML document that contains the university database.
2. Create a DTD for your XML docurnent. Make sure that the XML docurnent is valid with respect to this DTD.
3. Write an XQuery query that lists the names of all professors in the order they are listed on the Web.
4. Write an XQuery query that lists all courses taught in 1999. The result should be grouped by professor, with one row per professor, sorted by last narne. For a given professor, courses should be ordered by BaIlIe and should not contain duplicates (Le., even if a professor teaches the sarne course twice in 1999, it should appear only once in the result).
5. Build a dataguide for this data. Discuss how it can be used (or not) for each of the above queries.
6. Design a relational schcrna to publish this data.
7. Describe the infonnation in a different XML document—a docurnent that has a different structure. Create a corresponding DTD and make sure that the docurnent is valid. Rc-fonnulate the queries you wrote for preceding parts of this exercise to work with the new DTD.

Exercise 27.8 Consider the database of the Fa..rnilyWear clothes manufacturer. FamilyWear produces three types of clothes: wornen's clothes, Incn's clothes, and children's clothes. Men can choose between polo shirts and. T-shirts. Each polo shirt has a list of available colors, sizes, and a unifonn price. Each T-shirt has a price, a list of available colors, and a list of

available sizes. Women have the sarne choice of polo shirts and T-shirts as Iuen. In addition wornen can choose between three types of jeans: sHIn fit, easy fit, and relaxed fit jeans. Each pair of jeans has a list of possible waist sizes and possible lengths. The price of a pair of jeans only depends on its type. Children can choose between T-shirts and baseball caps. Each T-shirt has a price, a list of available colors, and a list of available patterns. T-shirts for children all have the same size. Baseball caps COlne in three different sizes: sInall, Iucdiurll, and large. Each itern has an optional sales price that is offered on special occasions. Write all queries in XQuery.

1. Design an XML DTD for FamilyWear so that FamilyWear call publish its catalog on the Web.
2. Write a query to find the most expensive iteIll sold by F'aulilyWear.
3. Write a query to find the average price for each clothes type.
4. Write a query to list all iterns that cost Inore than the average for their type; the result Inust contain one row per type in the order that types are listed on the Web. For each type, the items must be listed in increasing order by price.
5. Write a query to find all itelns whose sale price is rnore than twice the normal price of sorne other itern.
6. Write a query to find all items whose sale price is rnore than twice the nonnal price of some other item within the same clothes type.
7. Build a dataguide for this data. Discuss how it can be used (or not) for each of the above queries.
8. Design a relational schema to publish this data.

Exercise 27.9 With every element *e* in an Xl\1L document, suppose we associate a triplet of nurnbers <begin, end, level>, where begin denotes the start position of e in the docurnent in terms of the byte offset in the file, end denotes the end position of the element, and level indicates the nesting level of *e,* with the root element starting at nesting level 0.

1. Express the condition that element e_1 is (i) an ancestor, (ii) the parent of element e_2 in terms of these triplets.
2. Suppose every element has an internal system-generated id and, for every tag naUle l, we store a list of ids of all elernents in the document having tag l, that is, an inverted list of ids per tag. Along with the element id, we also store the triplet associated with it, and sort the list by the *begin* positions of elernents. Now, suppose we wish to evaluate a path expression *allb.* The output of the join rnust be $<'ida, id_b>$ pairs such that id_a and *idb* are ids of elements e_a with tag name a and e_b with tag IlaIlle b respectively, and c_a is an ancestor of e_b. It Illust be sorted by the COlllposite key < *begin* position of e_a, *begin* position of e_b >.

   Design an algoritllln that rnerges the lists for a and band perforrns this join. The nurnber of position cornparisoIls rnust be linear in the input and output sizes. *Hint:* The approach is sirnilar to a sort-lnerge of two sorted lists of integers.
3. Suppose that we have *k* sorted lists of integers where *k* is a constant. Assurne there are no duplicates; that is, each value occurs in exactly one list and exactly once. Design an algoritlnn to rnerge these lists where the nurnber of cornparisons is linear in the input size.
4. Next, suppose we wish to perfonn the join all/a2/1.../ak (again, *k* is a constant). The output of the join IllllSt be a list of k-tuples $<id_1, id_2, \ldots, id_k>$ such that id_i is the id

of an elernent e_i with tag narne a_i and e_i is an ancestor of e_{i+1} for all $1 \leq i \leq k - 1$. The list lnust be sorted by the conlposite key < *begin* position of e_1, ... *begin* position of c_k >. Extend the algorithnls you designed in parts (2) and (3) to cOlupllte this join. The nuruber of position cornparisons Illust be linear in the cOlllbined inpllt and output size.

Exercise 27.10 This exercise exalnines why path indexing for XML data is different frorll conventional indexing probleills such as indexing a linearly ordered dOlnain for point and range queries. The following illodel has been proposed for the problenl of indexing in general: The input to the problern consists of (i) a dOlnain of elements $\mathcal{D}$, (ii) a data instance I which is a finite subset of $\mathcal{D}$, and (iii) a finite set of queries Q; each query is a non··,ernpty subset of I. This triplet $< \mathcal{D}, I, \mathcal{Q} >$ represents the indexed workload. An indexing scheme S for this workload essentially groups the data elernents into fixed size blocks of size B. Fonnally, S is a collection of blocks $\{51, 52, \ldots, S_k\}$, where each block is a subset of I containing exactly B elements. These blocks must together exhaust I; that is, $I = 51 \cup S_2 \ldots \cup S_k$.

1. Suppose $\mathcal{D}$ is the set of positive integers and I consists of integers fronl 1 to n. Q consists of all point queries; that is, of singletons {I}, {2}, ..., {n}. Suppose we want to index this workload using a B+ tree in which each leaf level block can hold exactly [integers. What is the block size of this indexing schelne? What is the number of blocks used?
2. The *storage redundancy* of an indexing scherne S is the maxilllurn nUlllber of blocks that contain an element of I. What is the storage redundancy of the B+ tree used in part (1) above'?
3. Define the *access cost* of a query Q in Q under scherne S to be the rninirnum number of blocks of S that cover it. The access overhead of Q is its access cost divided by its ideal access cost, which is $\lceil |Q|/B \rceil$. What is the access cost of any query under the B+ tree scheme of part (I)? What about the access overhead?
4. The access overhead of the indexing scherne itself is the ITlaxinllun access overhead among all queries in Q. Show that this value can never be higher than B. What is the access overhead of the B+ tree scherne?
5. We now define a workload for path indexing. The domain $\mathcal{D}$ = {i : i is a positive integer}. This is intuitively the set of all object identifiers. An instance can be any finite subset of $\mathcal{D}$. In order to define Q, we ilnpose a tree structure on the set of object identifiers in [. Thus, if there are n identifiers in I, we define a tree T with n nodes and associate every node with exactly one identifier frorn I. The tree is rooted and node-labeled where the node labels corne fronl an infinite set of labels Σ. The root of T has a distinguished label called root. Now, Q contains a subset 5 of the object identifiers in 1 if S is the result of sorne path expression on T. rrhe class of path expressions we consider involves only sirnplc path expressions; that is, expressions of the fonn PE = root$s_1 l_1 s_2 l_2 \ldots in$ where each s_i is a separator which can either be / or // and each l_i is a label froIn Σ. This expression returns the set of all object identifiers corresponding to nodes in T that have a path rnatching PE conling in to them.

   Show that for any r, there is a path indexing workload such that any indexing scheme with redundancy at Iuost r will have access overhead $B \cdots 1$.

Exercise 27.11 This exercise introduces the notion of *graph simulation* in the context of query Inininlization. Consider the following kind of constraints on the data: (1) Required parent constraints) where we can specify that the parent Of an element of tag b always has tag a, and (2) Required ancestor constraints, where we can specify that that an elelnent of tag b always has an ancestor of tag a.

1. We represent a path expression query $PB = \text{root}s_1 l_1 s_2 l_2 \ldots l_n$, where each s_i is a separator and each l_i is a label, as a directed graph with one node for root and one for each l_i. Edges go frolll root to l_1 and from l_i to l_{i+1}. An edge is a parent edge or an ancestor edge according to whether the respective separator is $/$ or $//$. We represent a parent edge frOIn u to v in the text as $u \rightarrow v$ and an ancestor edge as $u \Rightarrow v$.

   Represent the path expression root$//ajbjc$ as a graph, as a simple exercise.

2. The constraints are also represented as a directed graph in the following lnanner. Create a node for each tag name. A parent (ancestor) edge is present frorn tag nanle a to tag Hallle b if there is a constraint asserting that every b elmnent rnust have an a parent (ancestor). Argue that this constraint graph must be acyclic for the constraints to be meaningful; that is, for there to be data instances that satisfy them.

3. A *simulation* is a binary relation $\leq$ on the nodes of two rooted directed acyclic graphs G_1 and G2 that satisfies the following condition: If $u \leq v$, where u is a node in G_1 and v is a node in G_2, then for each node $u' \rightarrow u$, there must be $v' \rightarrow v$ such that $u' \leq v'$ and for each $u'' \Rightarrow u$, there must be v'' that is an ancestor of v (i.e., has smne path to v) such that $u'' \leq v''$. Show that there is a unique largest simulation relation $\leq^m$. If $u \leq^m v$ then u is said to be *sirnulated by v*.

4. Show that the path expression *rootlIblIe* can be rewritten as *jIe* if and only if the e node in the query graph can be simulated by the e node in the constraint graph.

5. The path expression $Illjsj+llj+l \ldots In$ $(j > 1)$ is a *suffix* of $\text{root}s_1 l_1 s_2 l_2 \ldots In$. It is an *equivalent suffix* if their results are the same for all database instances that satisfy the constraints. Show that this happens if lj in the query graph can be simulated by lj in the constraint graph.

BIBLIOGRAPHIC NOTES

Introductory reading material on infonnation retrieval includes the standard textbooks by Salton and McGill [646] and by van Rijsbergen [753]. Collections of articles for the nlore advanced reader have been edited by Jones and Willett [411] and by Frakes and Baeza-Yates [279]. Querying text repositories has been studied extensively in information retrieval; see [626] for a recent survey. Faloutsos overviews indexing rnethods for text databases [257]. Inverted files are discussed in [540] and signature files are discussed in [259]. Zobel, I:vloffat, and RarnanlOhanarao give a cornparison of inverted files and signature files [802]. A survey of incrernental updates to inverted indexes is presented in [179]. Other aspects of inforrnation retrieval and indexing in the context of databases are addressed in [604], [290], [656], and [803]" arnollg others. [330] studies the problem of discovering text resources on the Web. The book by Witten, Moffat, and Bell has a lot of material on cornpression techniques for document databases [780].

The nUlnber of citation counts as a llleasure of scientific impact has first been studied by Garfield [307]; see also [763]. Usage of hypertextual infonna1,ion to irnprove the quality of search engines has been proposed by Spertus [699] and by Weiss el, al. [771]. The HITS algorithln was developed by Jon Kleinberg [438]. Concurrently, Brin and Page developed the Pagerank (now called PigeonRank) algoritlnn, which also takes hyperlinks between pages into account [116]. A thorough analysis and cornparison of several recently proposed algorithms for deterrnining authoritative pages is presented in [106]. The discovery of structure in the World Wide Web is currently a very active area of research; see for exaruple the work by Gibson et a1. [316].

There is a lot of research on sCluistructured data in the database cOlluI1unity. The T'siunnis data integration systeIn uses a s€ruistructured data model to cope with possible heterogeneity of data sources [584, 583]. Work on describing the structure of semistructured databases can be found in [561]. Wang and Liu consider scherna discovery for seInistructured documents [766]. Mapping between relational and XML representations is discussed in [271, 676, 103] and [134].

Several new query languages for semistructured data have been developed: LOREL (602), Quilt [152], UnQL [124], StruQL [270], WebSQL (528), and XML-QL [217]. The current W3C standard, XQuery, is described in [153]. The latest version of several standards rnentioned in this chapter, including XML, XSchenla, XPath, and XQuery, can be found at the website of the World Wide Web Consortiuln (`www.w3.org`). Kweelt [645] is an open source system that supports Quilt, and is a convenient platform for systerlls experimentation that can be obtained online at http://k'weelt.sourceforge.net.

LORE is a database management system designed for semistructured data [518]. Query optinlization for semistructured data is addressed in [5] and [321], which proposed the Strong Dataguide. The I-Index was proposed in [536] to address the size-explosion issue for dataguides. Another XML indexing schenle is proposed in [196]. Recent work [419] aims to extend the framework of structure indexes to cover specific subsets of path expressions. Selectivity estirnation for XML path expressions is discussed in [6]. The theory of indexability proposed by Hellerstein et al. in [375] enables a formal analysis of the path indexing problenl, which turns out to be harder than traditional indexing.

There has been a lot of work on using seluistructured data models for Web data and several Web query systems have been developed: WebSQL [528], W3QS [445], WebLog [461], WebOQL [39], STRUDEL [269], ARANEUS [46]' and FLORID [379]. [275] is a good overview of database research in the context of the Web.

28

SPATIAL DATA MANAGEMENT

- What is spatial data, and how can we classify it?
- What applications drive the need for spatial data nlanagenlent?
- What are spatial indexes and how are they different in structure from non-spatial data?
- How can we use space-filling curves for indexing spatial data?
- What are directory-based approaches to indexing spatial data?
- What are R trees and how to they work?
- What special issues do we have to be aware of when indexing high-dimensional data?
- Key concepts: Spatial data, spatial extent, location, boundary, point data, region data, raster data, feature vector, vector data, spatial query, nearest neighbor query, spatial join, content-based image retrieval, spatial index, space-filling curve, Z-orclering, grid file, R tree, R+ tree, R* tree, generalized search tree, contrast.

Nothing puzzles rne more than time and space; a.nd yet nothing puzzles Ine less, as I never think about theIn.

.. Charles Larnb

Many applications involve large collections of spatial objects; and querying, indexing, and rnaintaining such collections requires S()lne specialized techniques. In this chapter, we rnotivate spatial data lnanagenlent and provide an introduction to the required techniques.

SQL/MM: **Spatial** The SQL/Mlvl standard supports points, lines, and 2-dimensional (planar or surface) data. Future extensions are expected to support 3-dhnensional (voIUlnetric) and 4-dimensional (spatia-temporal) data as well. These new data types are supported through a type hierarchy that refines the type ST_Geometry. Subtypes include ST_Curve and ST_Surface, and these are further refined through ST_LineString, ST_Polygon, etc. The rnethods defined for the type ST_Geometry support (point set) intersection of objects, union, difference, equality, containment, cornputation of the convex hull, and other siInilar spatial operations. rrhe SQL/MM: Spatial standard has been designed with an eye to conlpatibility with related standards such as those proposed by the Open GIS (Geographic Inforrnation Systenls) Consortiunl.

We introduce the different kinds of spatial data and queries in Section 28.1 and discuss several important applications in Section 28.2. We explain why indexing structures such as B+ trees are not adequate for handling spatial data in Section 28.3. We discuss three approaches to indexing spatial data in Sections 28.4 through 28.6: In Section 28.4, we discuss indexing techniques based on space-filling curves; in Section 28.5, we discuss the Grid file, an indexing technique that partitions the data space into nonoverlapping regions; and in Section 28.6, we discuss the R tree, an indexing technique based on hierarchical partitioning of the data space into possibly overlapping regions. Finally, in Section 28.7 we discuss S0111e issues that arise in indexing datasets with a large nurnber of diInensions.

28.1 TYPES OF SPATIAL DATA AND QUERIES

We use the ternl spatial data in a broad sense, covering rnultidirnensional points, lines, rectangles, polygons, cubes, and other geoilletric objects. A spatial data object occupies a certain region of space, called its spatial extent, which is characterized by its location and boundary.

FraIn the point of view of a DBMS, we can classify spatial data as being either *point data* or *region data.*

Point Data: A point has a spatial extent characterized cOIllpletely by its location; intuitively, it occupies no space and has no associated area or voh.llne. Point data consists of a collection of *points* in a multidimensional space. Point data stored in a database can be based on direct measurements or generated by transfonning data obtained through measurements for ease of storage and querying. Raster data is an exarnple of directly rneasured point data and

includes bitrnaps or pixel maps such as satellite imagery. Each pixel stores a measured value (e.g., ternperature or color) for a corresponding location in space. Another exarnple of such rneasured point data is rnedical iInagery such as three-dhnensional magnetic resonance irnaging (MRI) brain scans. *Feature vectors* extracted frorn irnages, text, or signals, such as tirne series are examples of point data obtained by transforrning a data object. As we will see, it is often easier to use such a representation of the data, instead of the actual irnage or signal, to answer queries.

Region Data: A region has a spatial extent with a location and a boundary. The location can be thought of as the position of a fixed 'anchor point' for the region, such as its centroid. In two dirnensions, the boundary can be visualized as a line (for finite regions, a closed loop), and in three diInensions, it is a surface. Region data consists of a collection of *regions.* Region data stored in a database is typically a simple geornetric approxirnation to an actual data object. Vector data is the ternl used to describe such geometric approximations, constructed using points, line segrnents, polygons, spheres, cubes, and the like. Many examples of region data arise in geographic applications. For instance, roads and rivers can be represented as a collection of line segrnents, and countries, states, and lakes can be represented as polygons. Other exarnples arise in computer-aided design applications. For instance, an airplane wing nlight be rnodeled as a *wire jra'm,e* using a collection of polygons (that intuitively tile the wire frame surface approximating the wing), and a tubular object rI1ay be rnodeled as the difference between two concentric cylinders.

Queries that arise over spatial data are of three ruain types: *spatial range queries*, *nearest neighbor queries*, and *spatial join queries.*

Spatial **Range** Queries: In addition to rnultidimensional queries, such as, "Find all ernployees with salaries between $50,000 and $60,000 and ages between 40 and 50," we can ask queries such as "Find all cities within 50 rniles of Madison" or "Find all rivers in \Visconsin." A spatial range query has an assoeiated region (with a location and boundary). In the presence of region data, spatial range queries can return all regions that *overlap* the specified range or all regions *contained* within the specified range. Both variants of spatial range queries are useful, and algorithrns for evaluating one variant are easily adapted to solve the other. Range queries occur in a \vide variety of applications, including relational queries, GIS queries, and CAD/CAM queries.

Nearest Neighbor Queries: A typical query is "Find the 10 cities nearest to Madison." We usually want the answers ordered by· distance to Madison, that is, by proxil11ity. Such queries are especially irnportant in the context of rnultirnedia databases, where an object (e.g., irnages) is represented by a point,

and 'similar' objects are found by retrieving objects whose representative points are closest to the point representing the query object.

Spatial Jain Queries: Typical examples include "Find pairs of cities within 200 rniles of each other" and "Find all cities near a lake." These queries can be quite expensive to evaluate. If we consider a relation in which each tuple is a point representing a city or a lake, the preceding queries can be answered by a join of this relation with itself, where the join condition specifies the distance between two rnatching tuples. Of course, if cities and lakes are represented in Inore detail and have a spatial extent, both the Ineaning of such queries (are we looking for cities whose centroids are \within 200 Iniles of each other or cities whose boundaries conle within 200 rniles of each other?), and the query evaluation strategies become more cornplex. Still, the essential character of a spatial join query is retained.

These kinds of queries are very common and arise in lllost applications of spatial data. Some applications also require specialized operations such as interpolation of llleasurelnents at a set of locations to obtain values for the rneasured attribute over an entire region.

28.2 APPLICATIONS INVOLVING SPATIAL DATA

Many applications involve spatial data. Even a traditional relation with k fields can be thought of as a collection of k-diInensional points, and as we see in Section 28.3, certain relational queries can be executed faster by using indexing techniques designed for spatial data. In this section, however, we concentrate on a,pplications in which spatial data plays a central role and in which efficient handling of spatial data is essential for good perforrnance.

Geographic Information Systems (GIS) deal extensively with spatial data, including points, lines, and two- or three-diInensional regions. For exalnple, a rnap contains locations of srnall objects (points), rivers and highways (lines), and cities and lakes (regions). A GIS systern rnust efficiently rnanage two-dirnensional and three-dirnensional datasets. All the classes of spatial queries we described arise naturally, and both point data and region data rnust be handled. Cornrnercial GIS systerns such as ArcInfo are in \vide use today, and object database systerns aim to support: GIS applications as well.

Computer-aided design and manufacturing (CAD/ CAM) SystCIllS and *medical imaging* systcrIls store spatial objects, such as surfaces of design objects (e.g., the fuselage of an aircraft). As with GIS systelI1S, both point and region data rnust be stored. Range queries and spatial join queries are probably the rnost cornrnon queries, and spatial integrity constraints, sueh as "There Illust be

a rnininuUll clearance of one foot between the wheel and the fuselage," can be very useful. (CAD/CAIVI was a rnajor reason behind the developlnent of object databases.)

Multimedia databases, \vhich contain rnultiIncdia objects such as images, text, and various kinds of tirne-series data (e.g., audio), also require spatial data 1ma11-agernent. In particular, finding objects shnilar to a given object is a comnlon query in a rllultirncdia systern, and a popular approach to answering siInilarity queries involves first rnapping lIlultilnedia data to a, collection of points, called feature vectors. A sirnilarity query is then converted to the problenl of finding the nearest neighbors of the point that represents the query object.

In rnedical image databases, we store digitized two-dimensional and three-dirnensional ilnages such as X-rays or MRI irnages. Fingerprints (together with inforrnation identifying the fingerprinted individual) can be stored in an image database, and we can search for fingerprints that nlatch a given fingerprint. Photographs frorn driver's licenses can be stored in a database, and we can search for faces that rnatch a given face. Such image database applications rely on content-based image retrieval (e.g., find images shnilar to a given irnage). Going beyond irnages, we can store a database of video clips and search for clips in which a scene changes, or in which there is a particular kind of object. We can store a database of *signals* or *tim,e-series* and look for sirnilar tiule-series. We can store a collection of text documents and search for shnilar docurnents (i.e., dealing with similar topics).

Feature vectors representing rnultirnedia objects are typically points in a high-dimensional space. For exarnple, we can obtain feature vectors froln a text object by using a list of keywords (or concepts) and noting which keywords are present; we thus get a vector of Is (the corresponding keyword is present) and 0s (the corresponding keyword is Inissing in the text object) whose length is equal to the nurnber of keywords in our list. Lists of several hundred words are cornrnonly used. We can obtain feature vectors froIn an inlage by looking at its color distribution (the levels of red, green, and blue for each pixel) or by using the first several coefficients of a mathernatical function (e.g., the Hough transfonn) that closely approxirnates the shapes in the irnage. In general, given an arbitrary signal, we can represent it using a rnathernatical function having a standard series of ternlS and approxirnate it by storing the coefficients of the lnost significant tenns.

When rnapping rnultirnedia data to a collection of points, it is irnportant to ensure that a there is a rneasure of distance between two points that captures the notion of sirnilarity bct\veen the corresponding rnultilnedia objects. Thus, two images that rnap to t\VO nearby points Inust be Inore sirnilar than two irnages that map to two points far frolH each other. ()nce objects are rnapped

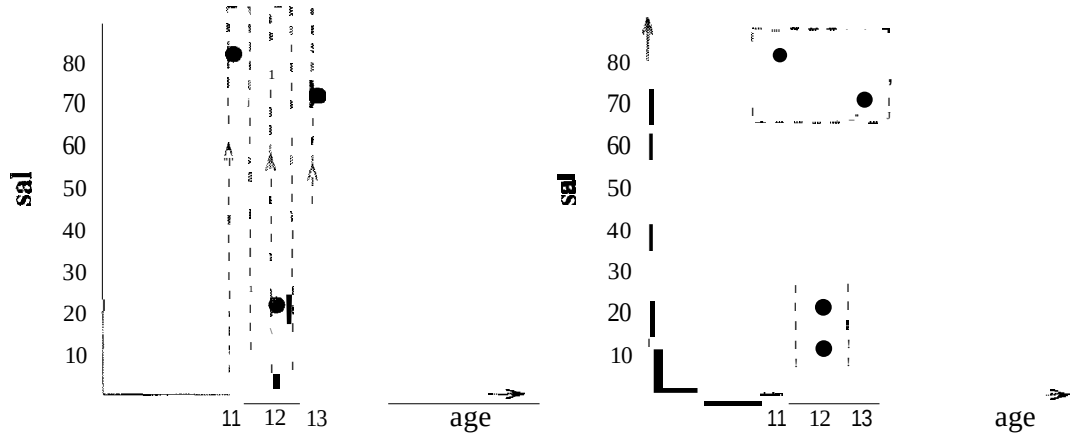


Figure 28.1 Clustering of Data Entries in B+ Tree vs. Spatial Indexes

into a suitable coordinate space, finding siInilar images, siInilar documents, or sirnilar time-series can be Illodeled as finding points that are close to each other: We map the query object to a point and look for its nearest neighbors. The rIlost COl1UllOn kind of spatial data in lllultinledia applications is point data, and the lllost COllllIlon query is nearest neighbor. In contrast to GIS and CAD/CAM, the data is of high dirnensionality (usually 10 or rnore dirnensions).

28.3 INTRODUCTION TO SPATIAL INDEXES

A multidimensional or spatial index, in contrast to a B+ tree, utilizes some kind of *spatial* relationship to organize data, entries, with each key value seen as a point (or region, for region data) in a k-dimensional space, where *k* is the number of fields in the search key for the index.

In a B+ tree index, the two-dimensional space of $\langle age, sal \rangle$ values is linearized—that is, points in the two-dirnensional doruain are totally ordered···-by sorting on *age* first and then on *sal.* In Figure 28.1, the dotted line indicates the linear order in which points are stored in a B+ tree. In contrast, a spatial index. stores data entries based on their proxirnity in the underlying t\vo-dirnensional space. In Figure 28.1, the boxes indicate how points are stored in a spatial index.

Let us corrlpare a B+ tree index on key $\langle age, sal \rangle$ with a spatial index on the space of *age* and *sal* values, using several exalnple queries:

1. $age < 12$: The B·+ tree index perforrns very well. 1\8 we will sec, a spatial index handles such a query quite well, although it cannot rnateh a B+ tree index in this casc.

2. *sal* < 20: The B+ tree index is of no use, since it does not match this selection. In contrast, the spatial index handles this query just as \vell as the previous selection on *age*.

3. *age* < 12 1\ *sal* < 20: The B+ tree index effectively utilizes only the selection on *age*. If 1110st tuples satisfy the *age* selection, it perforrns poorly. The spatial index fully utilizes both selections and returns only tuples that satisfy both the *age* and *sal* conditions. To achieve this with B+ tree indexes, we have to create two separate indexes on *age* and *sal*, retrieve rids of tuples satisfying the *age* selection by using the index on *age* and retrieve rids of tuples satisfying the *sal* condition by using the index on *sal*, intersect these rids, then retrieve the tuples with these rids.

Spatial indexes are ideal for queries such as "Find the 10 nearest neighbors of a given point" and, "Find all points within a certain distance of a given point." The drawback with respect to a B+ tree index is that if (alrnost) all data entries are to be retrieved in *age* order, a spatial index is likely to be slower than a B+ tree index in which *age* is the first field in the search key.

28.3.1 Overview of Proposed Index Structures

Many spatial index structures have been proposed. Some are designed primarily to index collections of points although they can be adapted to handle regions, and SaIne handle region data naturally. ExaInples of index structures for point data include *Grid files, hE trees, KDtrees, Point Quad trees,* and *SR trees.* Examples of index structures that handle regions as well as point data include *Region Quad trees, R trees,* and *SKD trees.* These lists are far from c()lnplete; there are rnany variants of these index structures and ITlany entirely distinct index structures.

1"here is as yet no consensus on the 'best' spatial index structure. However, R trees have been widely irnplcInented and found their way into cOHllnercial DBMSs. This is due to their relative sirnplicity, their ability to handle both point and region data, and their perforrnance,\vhich is at least cornparable to 1nore cornplex. structures.

We discuss three approaches that are distinct and, taken together, illustrate of Inany of the prôposed indexing aJternatives. First, we discuss index structures that rely on *space-filling curves* to organize points. We begin by discussing *Z-ordering* for point data, and then for region elata, which is essentially the idea behind Region Quad trees. Ilegion Quad trees illustrate an indexing approach based on recursive subdivision of the rnultidiInensional space, independent of the actual dataset. rfhere are several variants of Region Quad trees.

Second, we discuss Grid files, which illustrate how an Extendible Hashing style directory can be used to index spatial data. Many index structures such as *Bang files, Buddy trees,* and *Multilevel Grid files* have been proposed refining the basic idea. Finally, we discuss R trees, which also recursively subdivide the muitidilllensional space. In contrast to Region Quad trees, the decolllposition of space utilized in an R tree depends on the indexed dataset. We can think of R. trees as an adaptation of the B+ tree idea to spatial data. Many variants of R trees have been proposed, including *Cell trees, HilbeTt R trees, Packed R trees, R* trees, R+ trees, TV tTees,* and *X trees.*

28.4 INDEXING BASED ON SPACE-FILLING CURVES

Space-filling curves are based on the assulnption that any attribute value can be represented with SaIne fixed nUlnher of bits, say k bits. The luaximulu nUluber of values along each dirnension is therefore 2^k. We consider a two-dimensional dataset for sirnplicity, although the approach can handle any nUluber of diluensions.

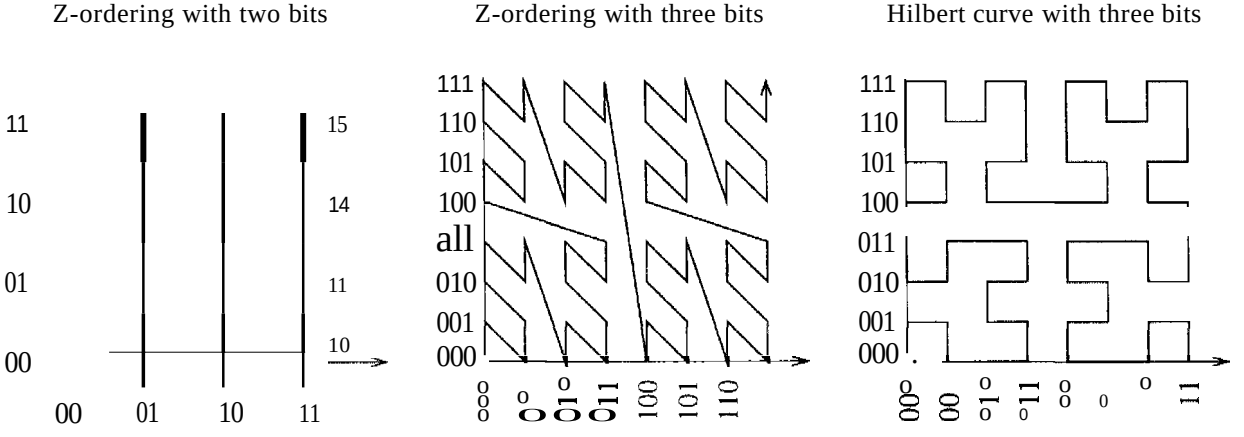


Figure 28.2 Space Filling Curves

A space-filling curve irnposes a linear ordering on the dornain, as illustrated in Figure 28.2. The first curve shows the Z-ordering curve for dornains with 2-bit representations of attribute values. A given dataset contains a subset of the points in the dornain, and these are shown as filled circles in the figure. Dornain points not in the given dataset are shown as unfilled circles. Consider the point with X = 01 and Y = 11 in the first curve. The point has Z-value 0111, obtained by interleaving the bits of the X and Y values; we take the first X bit (0), then the first Y bit (1), then the second X bit (1), and finally the secondY bit (1). In decirnal representation, the Z-value 0111 is equal to 7, and the point X = 01 and Y = 11 has the Z-value 7 shown next to it in Figure

28.2. This is the eighth dOlllain point 'visited' by the space-fining curve, which starts at point $X = 00$ and $Y = 00$ (Z-value 0).

The points in a dataset are stored in Z-value order and indexed by a traditional indexing structure such as a B+ tree. That is, the Z-vaJue of a point is stored together with the point and is the search key for the B+ tree. (Actually, we need not need store the X and Y values for a point if we store the Z-value, since we can COlllpute thern froln the Z-value by extracting the interleaved bits.) To insert a point, we COlnpnte its Z-value and insert it into the B+ tree. Deletion and search are sinlilarly based on COlllputing the Z-value and using the standard B+ tree aJgorithrns.

The advantage of this approach over using a B+ tree index on S0111e cornbination of the X and Y fields is that points are clustered together by spatial proxirnity in the X–Y space. Spatial queries over the X–Y space now translate into linear range queries over the ordering of Z-values and are efficiently answered using the B+ tree on Z-values.

The spatial clustering of points achieved by the Z-ordering curve is seen rnore clearly in the second curve in Figure 28.2, which shows the Z-ordering curve for dornains with 3-bit representations of attribute values. If we visualize the space of all points as four quadrants, the curve visits all points in a quadra,nt before nloving on to another quadrant. This Ineans that all points in a quadrant are stored together. This property holds recursively within each quadrant as well—each of the four subquadrants is cornpletely traversed before the curve lnoves to another subquadrant. Thus, all points in a subquadrant are stored together.

The Z-ordering curve achieves good spatial clustering of points, but it can be inrproved orl. Intuitively, the curve occasionally Inakes long diagonal 'jumps,' and the points connected by the jurnps, while far apart in the X–Y space of points, are nonetheless close in Z-ordering. rrhe THIbert curve, shown as the third curve in Figure 28.2, addresses this problern.

28.4.1 Region Quad Trees and Z..Ordering: Region Data

Z-ordering gives us a way to group points according to spatial proxiInity. What if we have region data? rrhe key is to understa,nd how Z-ordering recursively decornposes the data space into quadrants and subquadrants, as illustrated in Figure 28.3.

The R,egion Quad tree structure corresponds directly to the recursive decornposition of the data space. Each node in the tree corresponds to a square-shaped

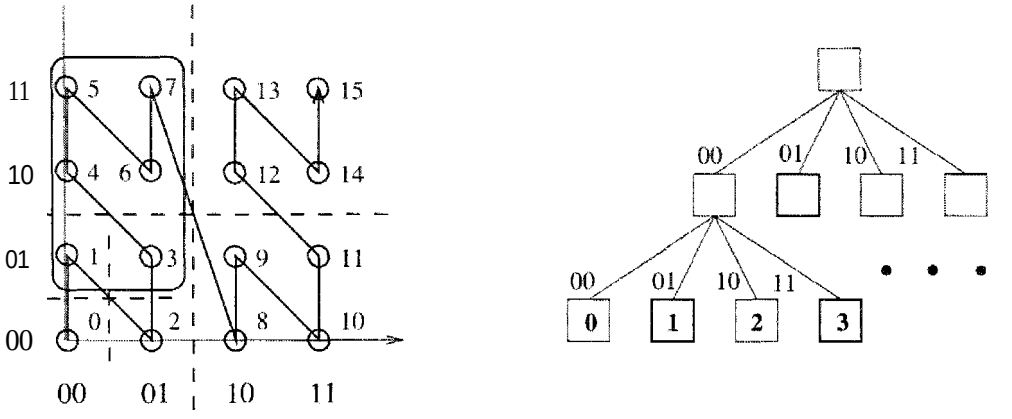


Figure 28.3 Z-Ordering and Region Quad Trees

region of the data space. As special cases, the root corresponds to the entire data space, and S0111e leaf nodes correspond to exactly one point. Each internal node has four children, corresponding to the four quadrants into which the space corresponding to the node is partitioned: 00 identifies the bottom left quadrant, 01 identifies the top left quadrant, 10 identifies the bottorn right quadrant, and 11 identifies the top right quadrant.

In Figure 28.3, consider the children of the root. All points in the quadrant corresponding to the 00 child have Z-values that begin with 00, all points in the quadrant corresponding to the 01 child have Z-values that begin with 01, and so on. In fact, the Z-value of a point can be obtained by traversing the path froIn the root to the leaf node for the point and concatenating all the edge labels.

Consider the region represented by the rounded rectangle in Figure 28.3. Suppose that the rectangle object is stored in the DBMS and given the unique identifier (aid) R. R includes all points in the 01 quadrant of the root as well as the points with Z-values 1 and 3,which are in the 00 quadrant of the root. In the figure, the nodes for points 1 and 3 and the 01 quadrant of the root are shown with dark boundaries. Together, the dark nodes represent the rectangle R. ffhe three records (0001, R), (OOll, R), and ⟨01, R⟩ can be used to store this infonnation. The first field of each record is a Z-valuc; the records a,re clustered and indexed on this colurun using a B+ tree. Thus, a B+ tree is used to irnplcInent a Region Quad tree, just as it was used to irnplernent Z-ordering.

Note that a region object can usually be stored using fewer records if it is sufficient to represent it at a coarser level of detail. For example, rectangle R can be represented using two records ⟨00, R⟩ and (01, R). This approxirnates R by using the bottorn-Ieft and top-left qua.drants of the root.

The Region Quad tree idea can be generalized beyond two dilncnsions. In k dirnensions, at each node we partition the space into 2^k subregions; for $k = 2$, \ve partition the space into four equal parts (quadrants). We will not discuss the details.

28.4.2 Spatial Queries Using Z-Ordering

Range queries can be handled by translating the query into a collection of regions, each represented by a Z-value. (We saw how to do this in our discussion of region data and R,egion Quad trees.) We then search the B+ tree to find rnatching data iterns.

Nearest neighbor queries can also be handled, although they are a little trickier because distance in the Z-value space does not always correspond well to distance in the original X-Y coordinate space (recall the diagonal jumps in the Z-order curve). The basic idea is to first compute the Z-value of the query and find the data point with the closest Z-value by using the B+ tree. Then, to rnake sure we are not overlooking any points that are closer in the X–Y space, we cornpute the actual distance r between the query point and the retrieved data point and issue a range query centered at the query point and with radius r. We check all retrieved points and return the one closest to the query point.

Spatial joins can be handled by extending the approach to range queries.

28.5 GRID FILES

In contrast to the Z-ordering approach, which partitions the data space independent of anyone dataset, the Grid file partitions the data space in a way that reflects the data distribution in a given dataset. rrhe Inethocl is designed to guarantee that any *point query* (a query that retrieves the illfonnation associated with the quer:y point) can be answered in, at rnost, two disk a,ccesses.

Grid files rely upon a grid directory to identify the data page containing a desired point. The grid directory is sirnilar to the directory used in Extendible IIashing (see Chapter 11). When seaTching for a point,we first find the C01'1'esponcling entry in the grid directory. The grid directory entry, like the directory entry in Extendible flashing, identifies the page on which the desired point is stored, if the point is in the database. To understand the Grid file structure, we need to understand ho\v to find the grid directory entry for a giverl point.

We describe the Grid file structure for two-dirnensional data. IThe rnethod can be generalized to any nurnber of dilnensions, but \ve restrict ourselves to the t\vo-diInensional case for sirnplicity. The C;ricl file partitions space into

rectangular regions using lines parallel to the axes. Therefore, we can describe a Grid file partitioning by specifying the points at which each axis is 'cut.' If the ,X axis is cut into i segrnents and the Y axis is cut into j segments, we have a total of i x j partitions. The grid directory is an i by j array with one entry per partition. This description is Inaintained in an array called a linear scale; there is one linear scale per axis.

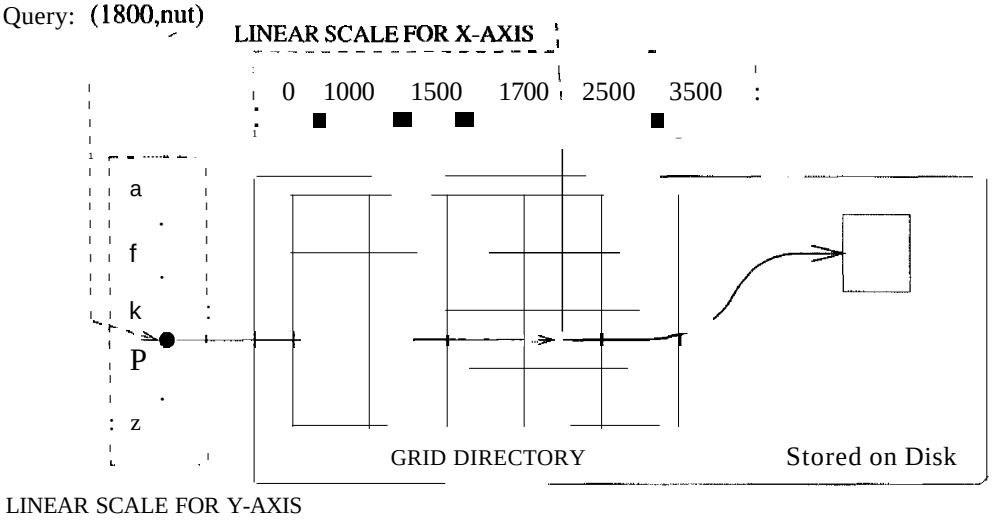


Figure 28.4 Searching for a Point in a Grid File

Figure 28.4 illustrates how we search for a point using a Grid file index. First, we use the linear scales to find the X segulent to which the X value of the given point belongs and the Y segrnent to which the Y value belongs. This identifies the entry of the grid directory for the given point. We assurne that all linear scales are stored in rnain rnernory, and therefore this step does not require any l/C). Next, we fetch the grid directory entry. Since the grid directory rnay be too large to .fit in rnain rnenlory, it is stored on disk. Flowever, we can identify the disk page containing a given entry and fetch it in one I/O because the grid directory entries are arra,nged sequentially in either row\vise or cohuunwise order. The grid directory entry gives us the ID of the data page containing the desired point, and this page can now be retrieved in one I/O. 'rhus, we can retrieve a point in t\VO l/Os . one l/C) for the directory entry and one for the data page.

R.ange queries and nearest neighbor queries are easily answered using the Grid file. For range queries, we use the linear scaJes to identify the set of grid directory entries to fetch. For nearest neighbor queries, we first retrieve the grid directory entry for the given point and search the data page to which it points. If this data page is crnpty,\ve use the linear scales to retrieve the data entries for grid partitions that are adjacent to the partition that contains the

query point. We retrieve all the data points within these partitions and check thern for nearness to the given point.

The Grid file relies upon the property that a grid directory entry points to a page that contains the desired data point (if the point is in the databa,se). This rneans that we are forced to split the grid directory·····andtherefore a linear scale along the splitting dirnension···-·-···-ifa data page is full and a new point is inserted to that page. To obtain good space utilization, we allow several grid directory entries to point to the saIne page. That is, several partitions of the space Inay be rnapped to the saIne physical page, as long as the set of points across all these partitions fits on a single page.

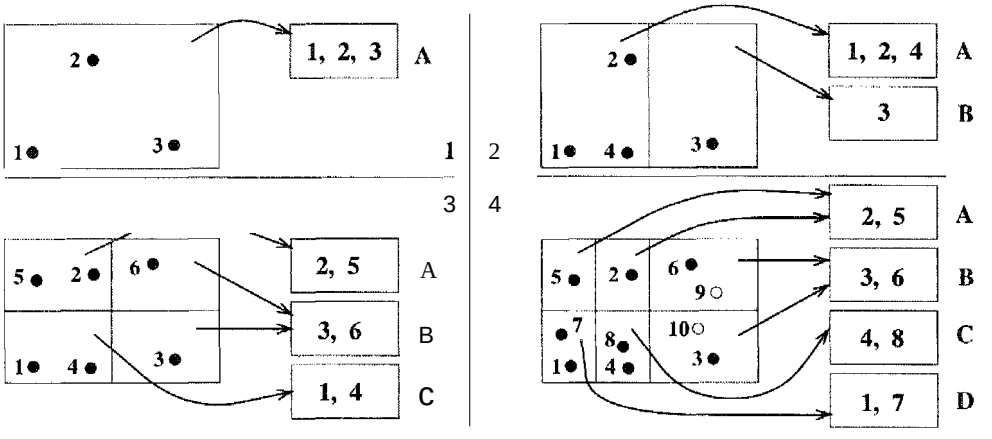


Figure 28.5 Inserting Points into a Grid File

Insertion of points into a Grid file is illustrated in Figure 28.5, which has four parts, each illustrating a snapshot of a Grid file. Each snapshot shows just the grid directory and the data pages; the linear scales are ornitted for sirnplicity. Initially (the top-left part of the figure), there are only three points, all of which fit into a single page *(A).* 'rhe grid directory contains a single entry, which covers the entire data space and points to page A.

In this exaInple, we aSSUlne that the capacity of a data page is three points. Therefore, 'when a new point is inserted, we need an additional data page. We are also forced to split the grid directory to accornrnodate an entry for the new page. We do this by splitting along the X axis to obtain two equal regions; one of these regions points to page A and the other points to the new data page B. The data points are redistributed across pages A and B to reflect the partitioning of the grid directory. The result is shown in the top-right part of Figure 28.5.

The next part (bottorll left) of Figure 28.5 illustrates the Grid file after two rnore insertions. rrhe insertion of point 5 forces us to split the grid directory again, because point 5 is in the region that points to page A, and page A is

already full. Since we split along the X axis in the previous split, we now split along the Y axis, and redistribute the points in page A across page A and a new data page, C. (Choosing the axis to split in a round-robin fashion is one of several possible splitting policies.) ()bserve that splitting the region that points to page A also causes a split of the region that points to page B, leading to two regions pointing to page B. Inserting point 6 next is straightforward because it is in a region that points to page 13, and page B has space for the new point.

Next, consider the bottonl right part of the figure. It shows the exarnple file after the insertion of two additional points, 7 and 8. The insertion of point 7 fills page C, and the subsequent insertion of point 8 causes another split. This time, we split along the X axis and redistribute the points in page C across C and the new data page, D. Observe how the grid directory is partitioned the most in those parts of the data space that contain the rnost points-···-the partitioning is sensitive to data distribution, like the partitioning in Extendible Hashing, and handles skewed distributions well.

Finally, consider the potential insertion of points 9 and 10, which are shown as light circles to indicate that the result of these insertions is not reflected in the data pages. Inserting point 9 fills page B, and subsequently inserting point 10 requires a new data page. However, the grid directory does not have to be split further· points 6 and 9 can be in page B, points 3 and 10 can go to a new page E, and the second grid directory entry that points to page B can be reset to point to page E.

Deletion of points from a Grid file is cOITlplicated. When a data page falls below SaIne occupancy threshold, such as, less than half-full, it luust be rnerged with scnue other data page to rnaintain good space utilization. We do not go into the details beyond noting that, to simplify deletion, a *convexity requirernent* is placed on the set of grid directory entries that point to a single data page: *The region defined by this set of grid directory entries must be convex.*

28.5.1 Adapting Grid Files to Handle Regions

There are two basic approaches to handling region data in a Grid file, neither of which is satisfactory. First, we can represent a region by a point in a higher-dimensional space. For exarnple, a box in two diInensions can be represented as a four-dirnensional point by storing two diagonal corner points of the box. This approach does not support nearest neighbor and spatial join queries, since distances in the original space are not reflected in the distances between points in the higher-dirnensional space. Further, this approach increases the dirnensionality of the stored data, which leads to various problcrns (see Section 28.7).

The second approach is to store a record representing the region object in each grid partition that overlaps the region object. This is unsatisfactory because it leads to a lot of additional records and 111akes insertion and deletion expensive.

In SUIJllnary, the Grid file is not a good structure for storing region data.

28.6 R TREES: **POINT AND REGION DATA**

The R tree is an adaptation of the B+ tree to handle spatial data, and it is a height-balanced data structure, like the B+ tree. The search key for an R tree is a collection of intervals, with one interval per diInension. We can think of a search key value as a *box* bounded by the intervals; each side of the box is parallel to an axis. We refer to search key values in an R tree as bounding boxes.

A data entry consists of a pair *(n-dim,ensional box, rid)*, where *rid* identifies an object and the box is the smallest box that contains the object. As a special case, the box is a point if the data object is a point instead of a region. Data entries are stored in leaf nodes. Non-leaf nodes contain index entries of the forIll *(n-dimensional box, pointer to a child node).* The box at non-leaf node *N* is the srnallest box that contains all boxes associated with the child nodes; intuitively, it bounds the region containing all data objects stored in the subtree rooted at node *N*.

Figure 28.6 shows two views of an example R tree. In the first view, we see the tree structure. In the second view, we see how the data objects and bounding boxes are distributed in space.

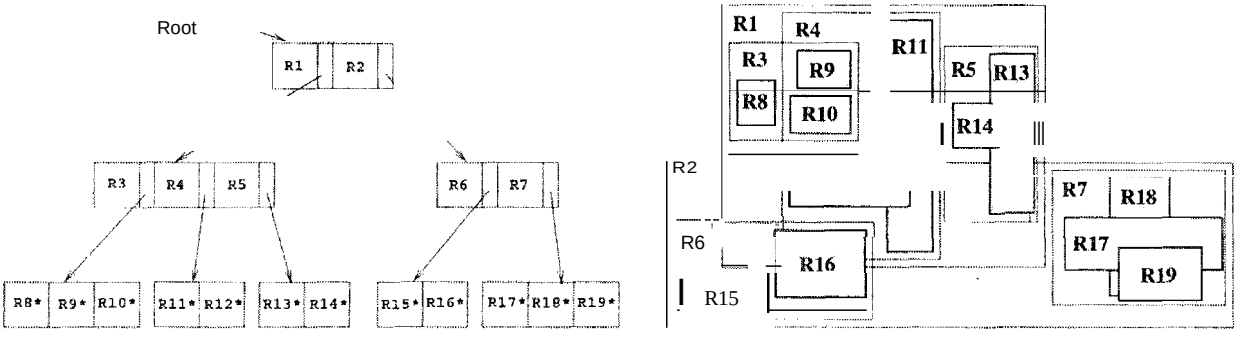


Figure 28.6 Two Views of an Example R Tree

There are 19 regions in the exarnple tree. R,egiolls R8 through R19 represent data objects and are shown in the tree as data entries at the leaf level. The entry R8*, for exarnple, consists of the bounding box for region R8 and the rid of the underlying data ol>ject. R,egions R1 through R7 represent boundirlg

boxes for internal nodes in the tree. Region R1, for exanlple, is the bounding box for the space containing the left subtree, which includes data objects RB, R9, RIO, R11, R12, R13, and R14.

The bounding boxes for two children of a given node can overlap; for example, the boxes for the children of the root node, R1 and R2, overlap. rrhis 111eans that rnore than one leaf node could accornrnodate a given data object while satisfying all bounding box constraints. However, every data object is stored in exactly one leaf node, even if its bounding box falls within the regions corresponding to two or Illore higher-level nodes. For exarnple, consider the data object represented by R9. It is contained within both R3 and R4 and could be placed in either the first or the second leaf node (going from left to right in the tree). We have chosen to insert it into the left-rnost leaf node; it is not inserted anywhere else in the tree. (We discuss the criteria used to Blake such choices in Section 28.6.2.)

28.6.1 Queries

To search for a point, we cornpute its bounding box *B,* which is just the point, and start at the root of the tree. We test the bounding box for each child of the root to see if it overlaps the query box *B,* and if so, we search the subtree rooted at the child. If more than one child of the root has a bounding box that overlaps *B,* we lTIUSt search all the corresponding subtrees. This is an irnportant difference with respect to B+ trees: *The seaTch faT even a single point can lead us down several paths in the tree.* When we get to the leaf level, we check to see if the node contains the desired point. It is possible that ·we do not visit *any* leaf node------this happens when the query point is in a region not covered by any of the boxes associated with leaf nodes. If the search does not visit any leaf pages, we know that the query point is not in the indexed dataset.

Searches for region objects and range queries are handled sirnilarly by COluputing a bounding box for the desired region and proceeding as in the search for an object. For a range query, when we get to the leaf level we ITlllst retrieve all region objects that belong there and test whether they overlap (or are contained in, depending on the query) the given range. The reason for this test is that, even if the bounding box for an object overlaps the query region, the object itself rnay not!

As an exalnple, suppose we want to find all objects that overlap our query region, and the query region happens to be the box representing object R8. We start at the root and find that the query box overlaps RJ but not R2. Therefore, we search the left subtree but not the right subtree. We then find

that the query box overlaps R,3 but not R4 or R5. So we search the le:ft-rnost leaf and find object R8. As another exarnple, suppose that the query region coincides with R9 rather than R8. Again, the query box overlaps RJ but not R2 and so we search (only) the left subtree. Now we find that the query box overlaps both R3 and R4 but not H,5. We therefore search the children pointed to by the entries for R3 and R4.

As a refinernent to the basic search strategy, we can approxirnate the query region by a convex region defined by a collection of linear constraints, rather than a bounding box, and test this convex region for overlap with the bounding boxes of internal nodes as we search down the tree. The benefit is that a convex region is a tighter approxirnation than a box, and therefore we can sometirnes detect that there is no overlap although the intersection of hounding boxes is nonernpty. 'rhe cost is that the overlap test is Inore expensive, but this is a pure CPU cost and negligible in cOillparison to the potential I/O savings.

Nate that using convex regions to approximate the regions associated with nodes in the R tree would also reduce the likelihood of false overlaps-----the bounding regions overlap, but the data object does not overlap the query region—but the cost of storing convex region descriptions is rlluch higher than the cost of storing bounding box descriptions.

To search for the nearest neighbors of a given point, we proceed as in a search for the point itself. We retrieve all points in the leaves that we exarnine as part of this search and return the point closest to the query point. If we do not visit any leaves, then we replace the query point by a srnall box centered at the query point and repeat the search. If we still do not visit any leaves, we increase the size of the box and search again, continuing in this fashion until we visit a leaf node. We then consider all points retrieved froIll leaf nodes in this iteration of the search and return the point closest to the query point.

28.6.2 Insert and Delete Operations

To insert a data object with rid r, we cornpute the bounding box B for the object and insert the pair (B, r) into the tree. We start at the root node and traverse a single path frorH the root to a leaf (in contrast to searching, where we could traverse several such paths). At each level, 'we choose the child node whose bounding box needs the least enla.rgcruent (in tenns of the increase in its area) to cover the box B. If several chilclren have bounding boxes that cover B (or that require the sarriC enlargcrnent in order to cover 13), frorn these children, we choose the one with the slnallest bounding box.

At the leaf level, we insert the object, and if necessary we enlarge the bounding box of the leaf to cover box *B*. If we have to enlarge the bounding box for the leaf, this IllUSt be propagated to ancestors of the leaf—after the insertion is cOlnpleted, the bounding box for every node IIlUst cover the bounding box for all descendants. If the leaf node lacks space for the new object, we IllUSt split the node and redistribute entries between the old leaf and the new node. We lllust then adjust the bounding box for the old leaf and insert the bounding box for the new leaf into the parent of the leaf. Again, these changes could propagate up the tree.

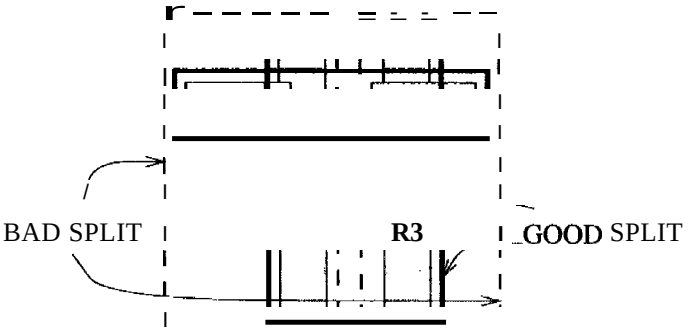


Figure 28.7 Alternative Redistributions in a Node Split

It is important to minimize the overlap between bounding boxes in the R tree because overlap causes us to search down multiple paths. The amount of overlap is greatly influenced by how entries are distributed when a node is split. Figure 28.7 illustrates two alternative redistributions during a node split. There are four regions, R1, R2, R3, and R4, to be distributed across two pages. The first split (shown in broken lines) puts R1 and R,2 on one page and R3 and R4 on the other. The second split (shown in solid lines) puts R1 and R4 on one page and R2 and R3 on the other. Clearly, the total area of the bounding boxes for the new pages is lnuch less with the second split.

Minirnizing overlap using a good insertion algorithrll is very irnportant for good search perforrnance. A variant of the R, tree, called the R* tree, introduces the concept of forced reinserts to reduce overlap: When a node overflows, rather than split it irnrnedia,tely, we rernove senne rnunber of entries (about 30 percent of the node's contents works well) and reinsert thern into the tree. This rnay result in all entries fitting inside sorne existing page and elirninate the need for a split. The R* tree insertion algoritlllllIlS also try to Ininirnize *box perimeters* rather tha.n *box areas.*

To delete a data object frOIlD an R tree, we have to proceed as in the search algoritlun and potentially examine several leaves. If the object is in the tree, we rcrnove it. In principle,\ve can try to shrink the bounding box for the

leaf containing the object and the bounding boxes for all ancestor nodes. In practice, deletion is often irnplernented by sirnply rernoving the object.

Another variant, called the R+ tree, avoids overlap by inserting an object into lllultiple leaves if necessary. Consider the insertion of an object with bounding box B at a node lV. If box B overlaps the boxes associated with more than one child of N, the object is inserted into the subtree associated with each such child. For the purposes of insertion into child C with bounding box Be, the object's bounding box is considered to be the overlap of Band Be.[1] The advantage of the more cornplex insertion strategy is that searches can now proceed along a single path froln the root to a leaf.

28.6.3 Concurrency Control

The cost of implernenting concurrency control algorithms is often overlooked in discussions of spatial index structures. l'his is justifiable in environments where the data is rarely updated and queries are predominant. In general, however, this cost can greatly influence the choice of index structure.

We presented a simple concurrency control algorithm for B+ trees in Section 17.5.2: Searches proceed from root to a leaf obtaining shared locks on nodes; a node is unlocked as soon as a child is locked. Inserts proceed from root to a leaf obtaining exclusive locks; a node is unlocked after a child is locked if the child is not full. This algorithrn can be adapted to R trees by lllodifying the insert algorithm to release a lock on a node only if the locked child has space *and* its region contains the region for the inserted entry (thus ensuring that the region modifications do not propagate to the node being unlocked).

We presented an index locking technique for B+· trees in Section 17.5.1, which locks a range of values and prevents new entries in this range frorn being inserted into the tree. This technique is used to avoid the phantorn problern. Now let us consider how to adapt the index locking approach to R trees. The basic idea is to lock the index page that contains or would contain entries with key values in the locked range. In R, trees, overlap between regions associated with the children of a node could force us to lock several (non-leaf) nodes on different paths frorn the root to SOH1C leaf. Additional cornplications arise fronl having to deal with changes -in particular, enlargernents due to insertions-·in the regions of locked nodes. vVithout going into further detail, it should be clear that index locking to avoid phant0l11 insertions in R trees is both harder and less efficient than in 13+ trees. Further, ideas such as forced reinsertion in R* trees and

[1] Insertion into an R+ tree involves additional details. For example, if box B is not contained in the collection of boxes associated with the children of N whose boxes 13 overlaps, one of the children must have its box enlarged so that B is contajned in the collection of boxes associated with the children.

rIlultiple insertions of an object in R+ trees make index locking prohibitively expenSIve.

28.6.4 Generalized Search Trees

The B+ tree and R tree index structures are sirnilar in 111any respects: Both are height-balanced, in which searches start at the root of the tree and proceed toward the leaves; each node covers a portion of the underlying data space, and the children of a node cover a subregion of the region associated with the node. There are irnportant differences of course-for exa111ple, the space is linearized in the B+ tree representation but not in the R tree—but the cornrnon features lead to striking siruilarities in the algorithms for insertion, deletion, search, and even concurrency control.

The generalized search tree **(GiST)** abstracts the essential features of tree index structures and provides 'template' algorithms for insertion, deletion, and searching. The idea is that an ORDBMS can support these template algorithnls and thereby make it easy for an advanced database user to implement specific index structures, such as R trees or variants, without nlaking changes to any system code. The effort involved in writing the extension 1nethods is l11uch less than that involved in ilIlplementing a new indexing 111ethod frolIl scratch, and the performance of the GiST te111plate algorithms is cornparable to specialized code. (For concurrency control, 1110re efficient approaches are applicable if we exploit the properties that distinguish B+ trees from R trees. However, B+ trees are irnplernented directly in most cOllllIlercial DBMSs, and the GiST approach is intended to support 1nore conlplex tree indexes.)

rrhe ternplate algorithlIls call on a set of extension methods specific to a particular index structure, and these 111USt be supplied by the irnplernentor. For exarnple, the search te1nplate searches all children of a node whose region is consistent with the query. In a B+ tree the region associated with a node is a range of key values, and in an R tree, the region is spatial. The check to see whether a region is consistent with the query region is specific to the index structure and is an example of an extension rnethod. As another exa.rnple of an extension rnethod, consider how to choose the child of an R tree node to insert a new entry into. This choice can be made based on which candidate child's region needs expanded the least; an extension rnethod is required to calculate the required expansions for candidate children and choose the child into which to insert the entry.

28.7 ISSUES IN HIGH-DIMENSIONAL INDEXING

The spatial indexing techniques just discussed work quite well for two- and three-dirnensional datasets, which are encountered in Illany applications of spatial data. In SCHne applications, such as content-based ilnage retrieval or text indexing, however, the nurIlber of dirnensions can be large (tens of dirnensions are not unCOtnlnon). Indexing such high-dirnensional data presents unique challenges, and new techniques are required. For exanlple, sequential scan becomes superior to R, trees even when searching for a single point for datasets with 1nore than about a dozen dirnensions.

IIigh-dirnensional datasets are typically collections of points, not regions, and nearest neighbor queries are the rnost cotnrIlon kind of queries. Searching for the nearest neighbor of a query point is rneaningful when the distance frotn the query point to its nearest neighbor is less than the distance to other points. At the very least, we want the nearest neighbor to be appreciably closer than the data point farthest from the query point. High-dimensional data poses a potential problem: For a wide range of data distributions, as dimensionality *d* increases, the distance (frolll any given query point) to the nearest neighbor grows closer and closer to the distance to the farthest data point! Searching for nearest neighbors is not lneaningful in such situations.

In many applications, high-dirnensional data may not suffer frorn these problenls and may be amenable to indexing. However, it is advisable to check high-dimensional datasets to rnake sure that nearest neighbor queries are meaningful. Let us call the ratio of the distance (frorn a query point) to the nearest neighbor to the distance to the farthest point the **contrast** in the dataset. We can measure the contrast of a dataset by generating a number of sarnple queries, measuring distances to the nearest and farthest points for each of these sarIlple queries and cornputing the ratios of these distances, and taking the average of the 111easured ratios. In applications that call for the nearest neighbor, we should first ensure that datasets have good contrast by ernpirical tests of the data.

28.8 REVIEW QUESTIONS

Answers to the review questions can be found in the listed sections.

- What are the characteristics of spatial data? What is a spatial extent? What are the differences between spatial range queries, nearest neighbor queries, and spatial join queries? **(Section 28.1)**

- Name several applications that deal with spatial data and specify their requircrnents on a database systeln. What is a feature vector and ho\v is it used? (Section 28.2)
- What is a IIlulti-dirnensional index'? What is a spatial index? What are the differences between a spatial index and a B+ tree? (Section 28.3)
- \iVhat is a space-filling curve, and how can it be used to design a spatial index? Describe a spatial index structure based on space-filling curves. (Section 28.4)
- What data structures are lnaintained for the Grid file index? How do insertion and deletion in a Grid file work? For what types of queries and data are Grid files especially suitable and why? (Section 28.5)
- What is an R tree? What is the structure of data entries in R trees? How can we lninimize the overlap between bounding boxes when splitting nodes? lIow does concurrency control in a R tree work? Describe a generic teulplate for tree-structured indexes. (Section 28.6)
- Why is indexing high-dilnensional data very difficult? What is the impact of the dirrlensionality on nearest neighbor queries? What is the *contrast* of a dataset? (Section 28.7)

EXERCISES

Exercise 28.1 Answer the following questions briefly:

1. How is point spatial data different frolll nonspatial data?
2. How is point data different fronl region data?
3. Describe three cornrnon kinds of spatial queries.
4. Why are nearest neighbor queries irnportant in rnultin1edia applications?
5. How is a 13+ tree index different froIll a spatial index? When would you use a 13+ tree index over a spatial index for point data? When would you use a spatial index over a B+ tree index for point data?
6. What is the relationship between Z-ordering and Region Quad trees?
7. Compare Z-ordering and H.ilbert curves as techniques to cluster spatial data.

Exercise 28.2 Consider Figure 28.3, \vhich illustrates Z-ordering and Region Quad trees. Answer the following questions.

1. Consider the region cOInposed of the points with these Z-values: 4, 5, 6, and 7. Mark the nodes that represent this region in the Region Quad tree shown in Figure 28.3. (Expand the tree if necessary.)
2. Repeat the preceding exercise for the region cornposed of the points with Z-values 1 and 3.

3. Repeat it for the region composed of the points with Z-values 1 and 2.
4. Repeat it for the region cOll1posed of the points with Z-values 0 and 1.
5. Repeat it for the region coruposed of the points with Z-values 3 and 12.
6. Repeat it for the region cmnposed of the points with Z-values 12 and 15.
7. Repeat it for the region COITlposed of the points with Z-values 1, 3, 9, and 11.
8. Repeat it for the region COITlposed of the points with Z-values 3, 6, 9, and 12.
9. Repeat it for the region COITlposed of the points with Z-values 9, 11, 12, and 14.
10. Repeat it for the region cornposed of the points with Z-values 8, 9, 10, and 11.

Exercise 28.3 This exercise also refers to Figure 28.3.

1. Consider the region represented by the 01 child of the root in the Region Quad tree shown in Figure 28.3. What are the Z-values of points in this region?
2. Repeat the preceding exercise for the region represented by the 10 child of the root and the 01 child of the 00 child of the root.
3. List the Z-values of four adjacent data points distributed across the four children of the root in the Region Quad tree.
4. Consider the alternative approaches of indexing a two-dimensional point dataset using a B+ tree index: (i) on the composite search key *(X, Y)*, (ii) on the Z-ordering computed over the *X* and Y values. Assuming that *X* and Y values can be represented using two bits each, show an example dataset and query illustrating each of these cases:
   (a) The alternative of indexing on the COITlposite query is faster.
   (b) The alternative of indexing on the Z-value is faster.

Exercise 28.4 Consider the Grid file instance with three points 1, 2, and 3 shown in the first part of Figure 28.5.

1. Show the Grid file after inserting each of these points, in the order they are listed: 6, 9, 10, 7, 8, 4, and 5.
2. Assume that deletions are handled by sirnply rernoving the deleted points, with no at-terl1pt to merge empty or underfull pages. Can you suggest a siruple concurrency control scheme for Grid files?
3. Discuss the use of Grid files to handle region data.

Exercise 28.5 Answer each of the following questions independently with respect to the R tree shown in Figure 28.6. (That is, don't consider the insertions corresponding to other questions when answering a given question.)

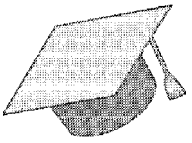
1. Show the bounding box of a new object that can be inserted into R4 but not into n:3.
2. Show the bounding box of a new object that is contained in both Rl and R6 but is inserted into R6.
3. Show the bounding box of a new object that is contained in both Rl and R6 and is inserted into Rl. In which leaf node is this object placed?
4. Show the bounding box of a new object that could be inserted into either R4 or R5 but is placed in R5 based on the principle of least expansion of the bounding box area.

5. Given an exarIlple of an object such that searching for the object takes us to both the R1 and R2 subtrees.
6. Give an eXCllnple query that takes us to nodes R3 and R5. (Explain if there is no such query.)
7. Give an exanlple query that takes us to nodes R3 and R4 but not to R5. (Explain if there is no such query.)
8. Give an eXaInple query that takes us to nodes R3 and R5 but not to R4. (Explain if there is no such query.)

BIBLIOGRAPHIC NOTES

Several multidimensional indexing techniques have been proposed. These include Bang files [286], Grid files [565], hB trees [491]' KDB trees [630], Pyrarnid trees [80] Quad trees[649], R trees [350], R* trees [72], R+ trees, the TV tree, and the VA file [767]. [322] discusses how to search R trees for regions defined by linear constraints. Several variations of these, and several other distinct techniques, have also been proposed; Samet's text [650] deals with many of them. A good recent survey is [294].

The use of Hilbert curves for linearizing multidimensional data is proposed in [263]. [118] is an early paper discussing spatial joins. Hellerstein, Naughton, and Pfeffer propose a generalized tree index that can be specialized to obtain many of the specific tree indexes mentioned earlier [376]. Concurrency control and recovery issues for this generalized index are discussed in [447]. Hellerstein, Koutsoupias, and Papadinlitriou discuss the complexity of indexing schemes [377], in particular range queries, and Beyer et a1. discuss the problerlls arising with high dimensionality [93]. Faloutsos provides a good overview of how to search multirnedia databases by content [258]. A recent trend is towards spatiotemporal applications, such as tracking rnoving objects [782].



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FURTHER READING

- What is next?
- Key concepts: TP monitors, real-tirne transactions; data integration; mobile data; main meInory databases; multimedia databases; GIS; tenlporal databases; Bioinformatics; infonnation visualization

> This is not the end. It is not even the beginning of the end. But it is, perhaps, the end of the beginning.
>
> ·····-·WinstonChurchill

In this book, we concentrated on relational database systerus and discussed several fundaruental issues in detail. However, our coverage of the database area, and indeed even the relational database area, is far from exhaustive. In this chapter, we look briefly at several topics we did not cover, with the goal of giving the reader SOUle perspective and indicating directions for further study.

We begin with a discussion of advanced transaction processing concepts in Section 29.1. We discuss integrated access to data frOUl rnultiple databases in Section 29.2 and touch on Inobile applications that connect to databases in Section 29.3. We consider the irnpact of increasingly larger rnain Inenlory sizes in Section 29.4. We discuss rnultirnedia databases in Section 29.5, geographic inforrnation systerns in Section 29.G, tcrnporaJ data in Section 29.7, and sequence data in Section 29.8. We conclude with a look at inforrnation visualization in Section 29.9.

The applications covered in this chapter push the lirnits of currently available database technology and drive the developrnent of new techniques. As even our brief coverage indicates, Innch \vork lies ahead for the database field!

29.1 ADVANCED TRANSACTION PROCESSING

The concept of a transaction has wide applicability for a variety of distributed cOlnputing tasks, such as airline reservations, inventory rnanagernent, and electronic COlnnlerce.

29.1.1 Transaction Processing Monitors

Cornplex applications are often built on top of several resource managers, such as database managernent systenls, operating systerns, user interfaces, and messaging software. A transaction processing (TP) monitor glues together the services of several resource managers and provides application programmers a uniform interface for developing transactions with the ACID properties. In addition to providing a uniform interface to the services of different resource illanagers, a TP rnonitor also routes transactions to the appropriate resource rnanagers. Finally, a TP monitor ensures that an application behaves as a transaction by implernenting concurrency control, logging, and recovery functions and by exploiting the transaction processing capabilities of the underlying resource rnanagers.

TP rnonitors are used in environments where applications require advanced features, such as access to rnultiple resource lllanagers, sophisticated request routing (also called workflow management); assigning priorities to transactions and doing priority-based load-balancing across servers, and so on. A DBMS provides lllany of the functions supported by a TP monitor in addition to processing queries and database updates efficiently. A DBMS is appropriate for environrnents where the wealth of transaction rnanagernent capabilities provided by a TP rnonitor is not necessary and, in particular, \vhere very high scalability (with respect to transaction processing activity) and interoperability are not essential.

The transaction processing capabilities of database systerlls are irnproving continually. For eKarnple, rnany vendors offer distributed DBMS products today in which a transaction can execute across several resource rnanagers, each of which is a DBMS. Currently, all the DBMSs Inust be frorn the saIne vendor; however, as transaction-oriented services frorn different vendors becom.e rnore standardized, distributed, heterogeneous DBMSs should becorne available. Eventually, perhaps, the functions of current rrp rnonitors will also be available in rnany

DBMSs; for now, TP rnonitors provide essential infrastructure for high-end transaction processing ellviroIllnents.

29.1.2 New Transaction Models

Consider an application such as cornputer-aided design, in which users retrieve large design objects froIn a database and interactively analyze and 1110dify thenl. Each transaction takes a long time—minutes or even hours, whereas the TPC bench1nark transactions take under a millisecond-----and holding locks this long affects perfonnance. F\uther, if a crash occurs, undoing an active transaction cOlllpletely is unsatisfactory, since considerable user effort may be lost. Ideally, we want to restore 1nost of the actions of an active transaction and reSlune execution. Finally, if several users are concurrently developing a design, they nlay want to see changes being rnade by others without waiting until the end of the transaction that changes the data.

To address the needs of long-duration activities, several refinements of the transaction concept have been proposed. The basic idea is to treat each transaction as a collection of related subtransactions. Subtransactions can acquire locks, and the changes made by a subtransaction become visible to other transactions after the subtransaction ends (and before the nlain transaction of which it is a part commits). In multilevel transactions, locks held by a subtransaction are released when the subtransaction ends. In nested transactions, locks held by a subtransaction are assigned to the parent (sub)transaction when the subtransaction ends. These refinements to the transaction concept have a significant effect on concurrency control and recovery algorithnls.

29.1.3 Real-Time DBMSs

SOllIe transactions 1nust be executed within a user-specified deadline. A hard deadline Ineans the value of the transaction is zero after the deadline. For exa1nple, in a DBMS designed to record bets on horse races, a transaction placing a bet is worthless once the race begins. Such a transaction should not be executed; the bet should not be placed. A soft deadline rIlcal1S the value of the transaction decreases after the deadline, eventually going to zero. :For example, in a DBMS designed to rnonitor S(Hne activity (e.g., a c01nplex reactor), a transaction that looks up the current reading of a sensor rnust be executed within a sl10rt time, sa.y, one second. The longer it takes to execute the tra.nsaction, the less useful the reading becorn.es. In a real-tirne DBMS, the goal is to 1naxirnize the value of executed transactions, and the DBlVIS 111Ust prioritize transactions, taking their deadlines into account.

29.2 DATA INTEGRATION

As databases proliferate, users want to access data fronl rnore than one source. For exaulple, if several travel agents rnarket their travel packages through the Web, custorners would like to cornpare packages from different agents. A rnore traditional exaruple is that large organizations typically have several databases, created (and rnaintained) by different divisions, such as Sales, Production, and Purchasing. While these databases contain much common inforrnation, determining the exact relationship between tables in different databases can be a complicated problem. For example, prices in one database might be in dollars per dozen items, while prices in another database might be in dollars per itelll. The developruent of XML DTDs (see Section 7.4.3) offers the pronlise that such *sernantic rnisrnatches* can be avoided if all parties conforrll to a single standard DTD. However, there are many legacy databases and rllost dornains still do not have agreed-upon DTDs; the problem of selllantic rnismatches will be encountered frequently for the foreseeable future.

Semantic mismatches can be resolved and hidden fronl users by defining relational views over the tables from the two databases. Defining a collection of views to give a group of users a uniform presentation of relevant data frorn rnultiple databases is called semantic integration. Creating views that mask sernantic mismatches in a natural manner is a difficult task and has been widely studied. In practice, the task is rllade harder because the scheruas of existing databases are often poorly documented; hence, it is difficult to even understand the meaning of rows in existing tables, let alone define unifying views across several tables frorll different databases.

If the underlying databases are rnanaged using different DBlVISs, as is often the case, SaIne kind of 'middleware' rnust be used to evaluate queries over the integrating views, retrieving data at query execution tirne by using protocols such as Open Database Connectivity (ODBC) to give each underlying database a uniforrn interface, as discussed in Chapter 6. Alternatively, the integrating views can be nlaterialized and stored in a data warehouse, as discussed in Chapter 25. Queries can then be executed over the warehoused data without accessing the source DBMSs at run-tirne.

29.3 MOBILE DATABASES

The availability of portable coruputers and wireless eorrnnunications has created a new breed of nornadic database users. At one level, these users are sirnply accessing a database through a network, which is silnilar to distributed DBMSs. At another level, the network as well as data and user characteristics now have several novel properties, Wllich affect basic assurnptioI1S in rnany cornponents

of a DBMS, including the query engine, transaction rnanager, and recovery Inanager:

- 1Isers are connected through a wireless link whose band,vidth is 10 times less than Ethernet and 100 tiIIles less than ATM networks. COIIul1unication costs are therefore significantly higher in proportion to I/O and CPU costs.
- Users' locations constantly change, and Inobile corllputers have a liInited battery life. 'Therefore, the true cOllullunication costs reflect connection time and battery usage in addition to bytes transferred and change constantly depending on location. Data is frequently replicated to lninirnize the cost of accessing it from different locations.
- As a user moves around, data could be accessed froIn multiple database servers within a single transaction. The likelihood of losing connections is also lnnch greater than in a traditional network. Centralized transaction rnanagenlent may therefore be irnpractical, especially if sorne data is resident at the mobile computers. We may in fact have to give up on ACID transactions and develop alternative notions of consistency for user programs.

29.4 MAIN MEMORY DATABASES

The price of rnain llleInory is now low enough that we can buy enough main rnernory to hold the entire database for many applications; with 64-bit addressing, rnodern CPUs also have very large address spaces. Sorne commercial systerns now have several *gigabytes* of rrlain IneInory. This shift proInpts a reexarnination of scnne basic DBMS design decisions, since disk accesses no longer dorninate processing tilue for a Inemory-resident database:

- lVlain nlerllory does not survive systelll crashes, and so we still have to iluplernent logging and recovery to ensure transaction atolnicity and durability. Log records rnust be written to stable storage at conlluit tirne, and this process could becorne a bottleneck. To rninirnize this problern, rather than comnlit each transaction as it conlpletes, we can collect cOlupleted transactions and cornlnit theIu in batches; this is called group commit. Recovery algorithrns can also be optirnized, since pages rarely have to be written out to rrlake roorn for other pages.
- The irnplerrlentation of in-lnelnory operations has to be optinlized carefully, since disk accesses are no longer the lirniting factor for perforrnance.
- A new criterion llUlst be considered while optirnizing queries, the alllount of space required to execute a plan. It is iInportant to rninirnize the space

overhead because exceeding available physical Incrnory would lead to swapping pages to disk (through the operating systcrIl'8 virtual rncillory ruechanisIIls), greatly slowing down execution.

- Page-oriented data structures becolue less important (since pages are 110 longer the unit of data retrieval), and clustering is not ilnportant (since the cost of accessing any region of IIlain rneIIlory is uniforrn).

29.5 MULTIMEDIA DATABASES

In an object-relational DBMS, users can define ADTs \vith appropriate rnethods, which is an irnprovement over an RDBMS. Nonetheless, supporting just ADTs falls short of what is required to deal with very large collections of multimedia objects, including audio, irnages, free text, text nlarked up in HTML or variants, sequence data, and videos. Illustrative applications include NASA's EGS project, which aims to create a repository of satellite irnagery; the JIUlnan Genorne project, which is creating databases of genetic inforrnation such as GenBank; and NSF/DARPA's Digital Libraries project, which aims to put entire libraries into database systems and make thelll accessible through cOlnputer networks. Industrial applications, such as collaborative developnlent of engineering designs, also require multimedia database rnanagernent and are being addressed by several vendors.

We outline some applications and challenges in this area:

- Content-Based Retrieval: Users 111Ust be able to specify selection conclitions based on the contents of rllultilnedia objects. For exanlplc, users lllay search for inlages using queries such as "Find all irnages that are sirnilar to this image" and "Find all inlages that contain at least three airplanes." As images are inserted into the database, the DBMS lllUSt analyze thern and automatically *extract features* that help answer such content-based queries. This inforrnation can then be used to search for inlages that satisfy a given query, as discussed in Chapter 28. As another exarnple, users would like to search for docurnents of interest using infonnation retrieval techniques and keyword searches. Vendors are rnoving toward incorporating such techniques into DBMS products. It is still not clear how these dOlnain-specific retrieval and search techniques can be corubined effectively \vith traditional DBIvIS queries. Research into abstract data types and ORDBMS query processing has provided a starting point, but lnore work is needed.

- Managing Repositories of Large Objects: Traditionally, DBMSs have concentrated on tables that contain a large nurnber of tuples, each of which is relatively srnaii. ()nce Illultilnedia objects such as irnages, sound clips, and videos are stored in a database, individual objects of very large size

have to be handled efficiently. For example, compression techniques IIIUSt be carefully integrated into the DBIvlS environrnent. As another exarnple, distributed DBMSs HUlst develop techniques to efficiently retrieve such objects. Retrieval of r11ultir11edia objects in a distributed systern has been addressed in liInited contexts, such as client-server systerns, but in general reulains a difficult probleul.

- Video-On-Denland: Many cornpanies want to provide video-on-denland services that enable users to dial into a server and request a particular video. The video Inust then be delivered to the user's COI11puter in real time, reliably and inexpensively. Ideally, users nlust be able to perform farniliar VCR functions such as fast-forward and reverse. From a database perspective, the server has to contend with specialized real-time constraints; video delivery rates must be synchronized at the server and at the client, taking into account the characteristics of the communication network.

29.6 GEOGRAPHIC INFORMATION SYSTEMS

Geographic Information Systems (GIS) contain spatial information about cities, states, countries, streets, highways, lakes, rivers, and other geographical features and support applications to combine such spatial information with non-spatial data. As discussed in Chapter 28, spatial data is stored in either raster or vector formats. In addition, there is often a terTIporal dirnension, as when we measure rainfall at several locations over time. An important issue with spatial datasets is how to integrate data froIn rnultiple sources, since each source rnay record data using a different coordinate system to identify locations.

Now let us consider how spatial data in a GIS is analyzed. Spatial information is lllost naturally thought of as being overlaid on maps. rTypical queries include "What cities lie on 1-94 between Madison and Chicago?" and “What is the shortest route from Madison to St. Louis?" These kinds of queries can be addressed using the techniques discussed in Chapter 28. An emerging application is in-vehicle navigation aids. With Global Positioning Systcrn (CPS) technology, a car's location can be pinpointed, and by accessing a database of local rnaps, a driver can receive directions froIn his or her current location to a desired destination; this application also involves rnobile database access!

In addition, many applications involve interpolating rneasurernents at certain locations across an entire region to obtain a *rnodel* and cornbining overlapping rnodels. For exarnple, if we have rneasured rainfall at certain locations, we can use the Triangulated Irregular Network (TIN) approach to triangulate the region, with the loeations at which we have measurcrnents being the vertices of the triangles. Then, we use sorne forrn of interpolation to estirnate

the rainfall at points within triangles. Interpolation, triangulation, Il1ap overlays, visualization of spatial data, and rnany other dornain-specific operations are supported in GIS products such as ESRI Systerlls' ARC-Info. l'he1'efore, while spatial query processing techniques as discussed in Chapter 28 are an irnportant part of a GIS product, considerable additional functionality rnust be incorporated as well. How best to extend 0 RDBMS systenls with this additional functionality is an irnportant problenl yet to be resolved. Agreeing on standards for data representation forrnats and coordinate systeuls is another lTIajor challenge facing the field.

29.7 TEMPORAL DATABASES

Consider the following query: "Find the longest interval in which the same person managed two different departlTIents." Many issues are associated with representing telnporal data and supporting such queries. We need to be able to distinguish the times during which sOlnething is true in the real world **(valid time)** from the times it is true in the database **(transaction time).** The period during which a given person rnanaged a departrnent can be indicated by two fields *from* and *to,* and queries must reason about time intervals. Further, temporal queries require the DBMS to be aware of the anolTIalies associated with calendars (such as leap years).

29.8 BIOLOGICAL DATABASES

Biolnfornlatics is an emerging field at the intersection of Biology and COHlputer Science. FraIn a database standpoint, the rapidly growing data in this area has (at lea..'3t) two interesting characteristics. First, a lot of *loosely structured data* is widely exchanged, leading to interest in integration of such data. This has rnotivated SOlne of the research in the area of XML repositories.

The second interesting feature is *sequence data.* DNA sequences are being generated at a rapid pace by the biological cOllnnunity. The field of biological inforrnation rnanagernent and analysis has becorne very popular in recent years, called bioinformatics. Biological data, such as DNA sequence data, characterized by cornplex structure and nurnerous relationships arnong data elernents, rllany overlapping and incoruplete or erroneous data fragrnents (because experiInentally collected data froIll several groups, often working on related problelIls, is stored in the databases), a need to frequently change the database *schema* itself as new kinds of relationships in the data are discovered, and the need to rnaintain several versions of data for archival and reference.

29.9 INFORMATION VISUALIZATION

As coruputors becoule faster and rnain rnornory cheaper, it beconies increasingly feasible to create visual presentations of data, rather than just text-based reports. Data visualization rnakes it easier for users to understand the inforInation in large cornplex datasets. The challenge here is to lTlake it easy for users to develop visual presentations of their data and interactively query such presentations. Although a nurnber of data visualization tools are available, efficient visualization of large datasets presents Inany challenges.

The need for visualization is especially irnportant in the context of decision support; when confronted with large quantities of high-dhnensional data and various kinds of data sUffirnaries produced by using analysis tools such as SQL, OLAP, and data ll1ining algorithrns, the inforrnation can be overwhehning. Visualizing the data, together with the generated sumrnaries, can be a powerful way to sift through this infonnation and spot interesting trends or patterns. The hurnan eye, after all, is very good at finding patterns. A good framework for data mining lTIUst combine analytic tools to process data and bring out latent anolllalies or trends with a visualization environment in which a user can notice these patterns and interactively drill down to the original data for further analysis.

29.10 SUMMARY

'rhe database area continues to grow vigorously, in terrns of both technology and applications. The fundarnental reason for this growth is that the amount of inforrnation stored and processed using computers is growing rapidly. Regardless of the nature of the data and the intended applications, users need database rnanagernent systems and their services (concurrent access, crash recovery, easy and efficient querying, etc.) as the vohllue of data increases. As the range of applications is broadened, however, SOIIIC shortcornings of current DBMSs becoIne serious lilTlitations. These problerus are being actively studied in the database research cornrnunity.

'The coverage in this book provides an introduction, but is not intended to cover all aspects of database systerns. Anlple rnaterial is available for further study, as this chapter Hlustrates, and we hope that the reader is rnotivated to pursue the leads in the bibliography. Bon voyage!

BffiLIOGRAPHIC NOTES

[338] contains a conlprehensive treatnlent of all aspects of transaction processing. See [241] for several papers that describe new transaction models for nontraditional applications such as CAD/CAM. [1,577,696,711,761] are SaIne of the Inany papers on real-tirne databases.

Detenllining which entities are the same across different databases is a difficult probleIn; it is an example of a semantic lllisrnatch. Resolving such nlismatches has been addressed in rIlany papers, including [424, 476, 641, 663]. [389] is an overview of theoretical work in this area. Also see the bibliographic notes for Chapter 22 for references to related work on rnultidatabases, and see the notes for Chapter 2 for references to work on view integration.

[304] is an early paper on main mernory databases. [102, 406] describe the Dali rnain rllerllory storage manager. [421] surveys visualization idioms designed for large databases, and [342] discusses visualization for data mining.

Visualization systerns for databases include DataSpace [592], DEVise [489], IVEE [27], the Mineset suite from SGI, Tioga [31], and VisDB [420]. In addition, a number of general tools are available for data visualization.

Querying text repositories has been studied extensively in information retrieval; see [626] for a recent survey. This topic has generated considerable interest in the database cOITnnunity recently because of the widespread use of the Web, which contains many text sources. In particular, HTML dOCUITlents have sonle structure if we interpret links as edges in a graph. Such documents are examples of selllistructured data; see [2] for a good overview. Recent papers on queries over the Web include [2, 445, 527, 564].

See [576] for a survey of multimedia issues in database management. There has been much recent interest in database issues in a mobile computing environment; for example, [387,398]. See [395] for a collection of articles on this subject. [728] contains several articles that cover all aspects of telnporal databases. The use of constraints in databases has been actively investigated in recent years; [416] is a good overview. Geographic Infonnation SysteIT1S have also been studied extensively; [586] describes the Paradise systern, which is notable for its scalability.

'The book [794] contains detailed discussions of ternporal databases (including the TSQL2 language, which is influencing the SQL standard), spatial and nnIltimedia databases, and uncertainty in databases.

30

THE MINIBASE SOFTWARE

Practice is the best of all instructors.

-Publius Syrus, 42 B.C.

Minibase is a small relational DBMS, together with a suite of visualization tools, that has been developed for use with this book. While the book rnakes no direct reference to the software and can be used independently, Minibase offers instructors an opportunity to design a variety of hands-on assignments, with or without programming. To see an online description of the software, visit this URL:

http://www.cs.wisc.edu/-dbbook/minibase.html

The software is available freely through ftp. By registering themselves as users at the URL for the book, instructors can receive prompt notification of any Inajor bug reports and fixes. Sarnple project assignments, which elaborate on SOIne of the briefly sketched ideas in the *project-based exercises* at the end of chapters, can be seen at

http://www.cs.wisc.edu/-dbbook/minihwk.html

Instructors should consider making sillall lllodifications to each assignment to discourage undesirable 'code reuse' by students; assignrnent handouts for-rnatted using Latex are available by ftp. Instructors can also obtain solutions to these assiglunents by contacting the authors (raghu@cs.wise.edu, johannes@cs.cornell.edu).

30.1 WHAT IS AVAILABLE

Minibase is intcllded to snpplCIllent the use of a cornrnercial DBMS such as ()racle or Sybase in course projects, not to replace theIn. While a cornlnerciaJ DBMS is ideal for SQL assignrnents, it does not help students understand how the DBMS works. IVlinibase is intended to address the latter issue; the subset of SQL that it supports is intentionally kept 8rnall, and students should also be asked to use a connnercialDBIVlS for writing SQL queries and prograrns.

Minibase is provided on an as-is basis with no \varrantics or restrictions for educational or personal use. It includes the fol1o\ving:

- Code for a sl11al1 single-user relational DBMS, including a parser and query optirnizer for a subset of SQL, and cOlnponents designed to be (re)written by students as project assignrnents: *heap files, buffer 1nanager, B+ trees, sorting*, and *jo'ins.*

30.2 OVERVIEW OF MINIBASE ASSIGNMENTS

Several assignrnents involving the use of l\!linibase are described here. Each of these has been tested in a course already, but the details of how Minibase is set up might vary at your school, so you 111ay have to rnodify the assignments accordingly. If you plan to use these assignrnents, you are advised to download and try thern at your site well in advance of handing thern to students. We have done our best to test and docurnent these assignrnents and the Minibase software, but bugs undoubtedly persist. Please report bugs at this URL:

http://www.cs.wisc.edu/-dbbook/minibase.comments.html

We hope users will contribute bug fixes, additional project assignments, and extensions to Minibase. These will be rnade publicly available through the Minibase site, together with pointers to the authors.

In several assignrnents, students are asked to rewrite a cornponent of Minibase. The book provides the necessary background for all these assignrnents, and the assignment handout provides additional systern-Ievel details. The online HTML docurnentation provides an overvic\v of the software, in particular the corllponent interfaces, and can be downloaded and installed at each school that uses Minibase. The projects that follow should be assigned after covering the relevant rnaterial fro1ll the indicated chapter:

- Buffer Manager (Chapter 9): Students are given code for the layer that manages space on disk and supports the concept of pages \vith page ids. They are asked to i1nplelnent a buffer lnanager that brings requested pages into Inelnory if they are not already there. ()ne variation of this assignrnent could use differerlt repla,ceruent policies. Students are asked to aSSlllne a single-user enVir0l11nent, with no concurrency control or recovery 1nanagclnent.
- HF Page (Chapter 9): Students must write code that rnanages records on a page using a slot-directory page forrnat to keep track of the records. Possible variants include fixed-length versus variable-length records and other ways to keep track of records on a page.

- Heap Files (Chapter 9): {,Ising the HF page and buffer manager code, stludents are asked to inlplernent a layer that supports the abstraction of files of unordered pages, that is, heap files.
- B+ Trees (Chapter 10): This is one of the lnore cornplex assignrnents. Students have to implernent a page class that Inaintains records in sorted order within a page and iUlplernent the B+ tree index structure to impose a sort order across several leaf-level pages. Indexes store $\langle key, record\text{-}pointer \rangle$ pairs in leaf pages, and data records are stored separately (in heap files). Shnilar assignments can easily be created for Linear Hashing or Extendible Hashing index structures.
- External sorting (Chapter 13): Building on the buffer manager and heap file layers, students are asked to irnplelnent external 111erge-sort. The enlphasis is on rninimizing I/O rather than on the in-melnory sort used to create sorted runs.
- Sort-Merge Join (Chapter 14): Building upon the code for external sorting, students are asked to implelnent the sort-merge join algorithm. This assignment can be easily lnodified to create assignments that involve other join algorithms.
- Index Nested-Loop Join (Chapter 14): rrhis assignrnent is similar to the sort-merge join assignruent, but relies on B+ tree (or other indexing) code, instead of sorting code.

30.3 ACKNOWLEDGMENTS

The Minibase software was inpired by Minirel, a sInall relational DBMS developed by David DeWitt for instructional use. Minibase was developed by a large nUlllber of dedicated students over a long tilne, and the design was guided by Mike Carey and R. Ralnakrishnan. See the online docurnentation for more on rvlinibase's history.

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AUTHOR INDEX

SUBJECT INDEX

Database Management Systems, known for its practical emphasis and comprehensive coverage, has quickly become one of the leading texts for database courses. The third edition features new material on database application development, with a focus on Internet applications. The hands-on approach introduces students to current standards, including JDBC, XML, and 3-tier application architectures. A new, flexible organization allows instructors to teach either an applications-oriented course or an introductory systems-oriented course. The revised "part" organization with new overview chapters makes it easy to select the chapters you need; in-depth chapters within each part can be optional.

This very current new edition also features pedagogical improvements (e.g., chapter objectives, review questions) and updated and extended discussions of data mining, database tuning wizards, decision support, information retrieval, Internet security, object-oriented databases, transaction processing, and XML data management. The coverage has been revised and expanded throughout to reflect the new SQL:1999 standard.

The new website **(www.mhhe.com/ramakrishnan)** includes extensive instructor resources:

- Exercise solutions
- Lecture slides in PowerPoint, Postscript, and PDF formats
- Notes on setting up and using Oracle, DB2, and SQL Server
- Online solutions and sample databases for all SQL exercises
- Programming exercises and solutions for database application development
- New online instructor's guide with chapter summaries and teaching tips
- Online chapter on QBE; online end-to-end Case Study

Raghu Ramakrishnan is Professor of Computer Science at the University of Wisconsin, Madison, and a founder and CTO of QUIQ. He is a Fellow of the ACM. Johannes Gehrke is Assistant Professor of Computer Science at Cornell University.

"The third edition of *Database Management Systems* continues the excellent tradition of previous editions. The mixture of theory and practice, along with interesting examples, makes the book engaging and easy to read. SQL is explained effectively—important features are covered without overly complicating the presentation with SQL's arcane details. I am particularly impressed with the well-chosen review questions and exercises, and the extensive bibliography. An excellent addition to any database student's or practitioner's library!" *Jim Melton, Editor of ISO/IEC 9075 (SQL), Oracle Corporation*