chapter 21

Introduction to Transaction Processing Concepts and Theory

The concept of transaction provides a mechanism for describing logical units of database processing. **Transaction processing systems** are systems with large databases and hundreds of concurrent users executing database transactions. Examples of such systems include airline reservations, banking, credit card processing, online retail purchasing, stock markets, supermarket checkouts, and many other applications. These systems require high availability and fast response time for hundreds of concurrent users. In this chapter we present the concepts that are needed in transaction processing systems. We define the concept of a transaction, which is used to represent a logical unit of database processing that must be completed in its entirety to ensure correctness. A transaction is typically implemented by a computer program, which includes database commands such as retrievals, insertions, deletions, and updates. We introduced some of the basic techniques for database programming in Chapters 13 and 14.

In this chapter, we focus on the basic concepts and theory that are needed to ensure the correct executions of transactions. We discuss the concurrency control problem, which occurs when multiple transactions submitted by various users interfere with one another in a way that produces incorrect results. We also discuss the problems that can occur when transactions fail, and how the database system can recover from various types of failures.

This chapter is organized as follows. Section 21.1 informally discusses why concurrency control and recovery are necessary in a database system. Section 21.2 defines the term *transaction* and discusses additional concepts related to transaction processing in database systems. Section 21.3 presents the important properties of atomicity, consistency preservation, isolation, and durability or permanency—called the

ACID properties—that are considered desirable in transaction processing systems. Section 21.4 introduces the concept of schedules (or histories) of executing transactions and characterizes the *recoverability* of schedules. Section 21.5 discusses the notion of *serializability* of concurrent transaction execution, which can be used to define correct execution sequences (or schedules) of concurrent transactions. In Section 21.6, we present some of the commands that support the transaction concept in SQL. Section 21.7 summarizes the chapter.

The two following chapters continue with more details on the actual methods and techniques used to support transaction processing. Chapter 22 gives an overview of the basic concurrency control protocols and Chapter 23 introduces recovery techniques.

21.1 Introduction to Transaction Processing

In this section we discuss the concepts of concurrent execution of transactions and recovery from transaction failures. Section 21.1.1 compares single-user and multiuser database systems and demonstrates how concurrent execution of transactions can take place in multiuser systems. Section 21.1.2 defines the concept of transaction and presents a simple model of transaction execution based on read and write database operations. This model is used as the basis for defining and formalizing concurrency control and recovery concepts. Section 21.1.3 uses informal examples to show why concurrency control techniques are needed in multiuser systems. Finally, Section 21.1.4 discusses why techniques are needed to handle recovery from system and transaction failures by discussing the different ways in which transactions can fail while executing.

21.1.1 Single-User versus Multiuser Systems

One criterion for classifying a database system is according to the number of users who can use the system **concurrently**. A DBMS is **single-user** if at most one user at a time can use the system, and it is **multiuser** if many users can use the system—and hence access the database—concurrently. Single-user DBMSs are mostly restricted to personal computer systems; most other DBMSs are multiuser. For example, an airline reservations system is used by hundreds of travel agents and reservation clerks concurrently. Database systems used in banks, insurance agencies, stock exchanges, supermarkets, and many other applications are multiuser systems. In these systems, hundreds or thousands of users are typically operating on the database by submitting transactions concurrently to the system.

Multiple users can access databases—and use computer systems—simultaneously because of the concept of **multiprogramming**, which allows the operating system of the computer to execute multiple programs—or **processes**—at the same time. A single central processing unit (CPU) can only execute at most one process at a time. However, **multiprogramming operating systems** execute some commands from one process, then suspend that process and execute some commands from the next

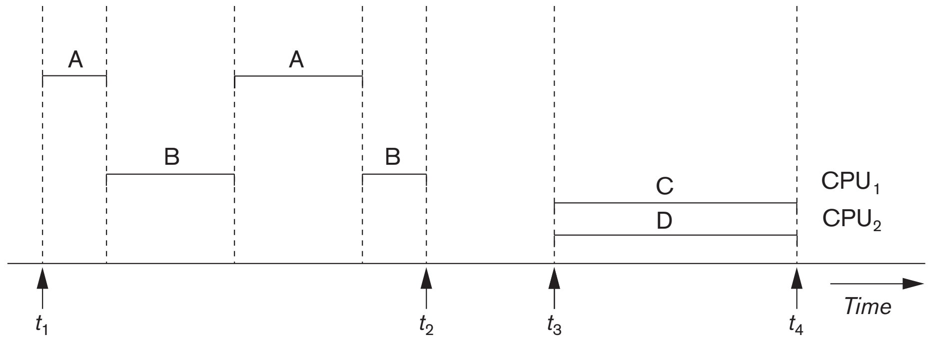


Figure 21.1
Interleaved processing versus parallel processing of concurrent transactions.

process, and so on. A process is resumed at the point where it was suspended whenever it gets its turn to use the CPU again. Hence, concurrent execution of processes is actually **interleaved**, as illustrated in Figure 21.1, which shows two processes, A and B, executing concurrently in an interleaved fashion. Interleaving keeps the CPU busy when a process requires an input or output (I/O) operation, such as reading a block from disk. The CPU is switched to execute another process rather than remaining idle during I/O time. Interleaving also prevents a long process from delaying other processes.

If the computer system has multiple hardware processors (CPUs), **parallel processing** of multiple processes is possible, as illustrated by processes C and D in Figure 21.1. Most of the theory concerning concurrency control in databases is developed in terms of **interleaved concurrency**, so for the remainder of this chapter we assume this model. In a multiuser DBMS, the stored data items are the primary resources that may be accessed concurrently by interactive users or application programs, which are constantly retrieving information from and modifying the database.

21.1.2 Transactions, Database Items, Read and Write Operations, and DBMS Buffers

A **transaction** is an executing program that forms a logical unit of database processing. A transaction includes one or more database access operations—these can include insertion, deletion, modification, or retrieval operations. The database operations that form a transaction can either be embedded within an application program or they can be specified interactively via a high-level query language such as SQL. One way of specifying the transaction boundaries is by specifying explicit **begin transaction** and **end transaction** statements in an application program; in this case, all database access operations between the two are considered as forming one transaction. A single application program may contain more than one transaction if it contains several transaction boundaries. If the database operations in a transaction do not update the database but only retrieve data, the transaction is called a **read-only transaction**; otherwise it is known as a **read-write transaction**.

The *database model* that is used to present transaction processing concepts is quite simple when compared to the data models that we discussed earlier in the book, such as the relational model or the object model. A **database** is basically represented as a collection of *named data items.* The size of a data item is called its **granularity**. A **data item** can be a *database record*, but it can also be a larger unit such as a whole *disk block*, or even a smaller unit such as an individual *field (attribute) value* of some record in the database. The transaction processing concepts we discuss are independent of the data item granularity (size) and apply to data items in general. Each data item has a *unique name*, but this name is not typically used by the programmer; rather, it is just a means to *uniquely identify each data item*. For example, if the data item granularity is one disk block, then the disk block address can be used as the data item name. Using this simplified database model, the basic database access operations that a transaction can include are as follows:

- **read_item(*X*).** Reads a database item named *X* into a program variable. To simplify our notation, we assume that *the program variable is also named X.*
- **write_item(*X*).** Writes the value of program variable *X* into the database item named *X*.

As we discussed in Chapter 17, the basic unit of data transfer from disk to main memory is one block. Executing a read_item(*X*) command includes the following steps:

1. Find the address of the disk block that contains item *X*.
2. Copy that disk block into a buffer in main memory (if that disk block is not already in some main memory buffer).
3. Copy item *X* from the buffer to the program variable named *X*.

Executing a write_item(*X*) command includes the following steps:

1. Find the address of the disk block that contains item *X*.
2. Copy that disk block into a buffer in main memory (if that disk block is not already in some main memory buffer).
3. Copy item *X* from the program variable named *X* into its correct location in the buffer.
4. Store the updated block from the buffer back to disk (either immediately or at some later point in time).

It is step 4 that actually updates the database on disk. In some cases the buffer is not immediately stored to disk, in case additional changes are to be made to the buffer. Usually, the decision about when to store a modified disk block whose contents are in a main memory buffer is handled by the recovery manager of the DBMS in cooperation with the underlying operating system. The DBMS will maintain in the **database cache** a number of **data buffers** in main memory. Each buffer typically holds the contents of one database disk block, which contains some of the database items being processed. When these buffers are all occupied, and additional database disk blocks must be copied into memory, some buffer replacement policy is used to choose which of the current buffers is to be replaced. If the chosen buffer has been modified, it must be written back to disk before it is reused.[1]

A transaction includes read_item and write_item operations to access and update the database. Figure 21.2 shows examples of two very simple transactions. The **read-set** of a transaction is the set of all items that the transaction reads, and the **write-set** is the set of all items that the transaction writes. For example, the read-set of T_1 in Figure 21.2 is $\{X, Y\}$ and its write-set is also $\{X, Y\}$.

Concurrency control and recovery mechanisms are mainly concerned with the database commands in a transaction. Transactions submitted by the various users may execute concurrently and may access and update the same database items. If this concurrent execution is *uncontrolled*, it may lead to problems, such as an inconsistent database. In the next section we informally introduce some of the problems that may occur.

21.1.3 Why Concurrency Control Is Needed

Several problems can occur when concurrent transactions execute in an uncontrolled manner. We illustrate some of these problems by referring to a much simplified airline reservations database in which a record is stored for each airline flight. Each record includes the *number of reserved seats* on that flight as a *named (uniquely identifiable) data item*, among other information. Figure 21.2(a) shows a transaction T_1 that *transfers* N reservations from one flight whose number of reserved seats is stored in the database item named X to another flight whose number of reserved seats is stored in the database item named Y. Figure 21.2(b) shows a simpler transaction T_2 that just *reserves* M seats on the first flight (X) referenced in transaction T_1.[2] To simplify our example, we do not show additional portions of the transactions, such as checking whether a flight has enough seats available before reserving additional seats.

(a)

T_1

read_item(X);
$X := X - N$;
write_item(X);
read_item(Y);
$Y := Y + N$;
write_item(Y);

(b)

T_2

read_item(X);
$X := X + M$;
write_item(X);

Figure 21.2
Two sample transactions. (a) Transaction T_1. (b) Transaction T_2.

[1] We will not discuss buffer replacement policies here because they are typically discussed in operating systems textbooks.

[2] A similar, more commonly used example assumes a bank database, with one transaction doing a transfer of funds from account X to account Y and the other transaction doing a deposit to account X.

When a database access program is written, it has the flight number, flight date, and the number of seats to be booked as parameters; hence, the same program can be used to execute *many different transactions*, each with a different flight number, date, and number of seats to be booked. For concurrency control purposes, a transaction is a *particular execution* of a program on a specific date, flight, and number of seats. In Figure 21.2(a) and (b), the transactions T_1 and T_2 are *specific executions* of the programs that refer to the specific flights whose numbers of seats are stored in data items X and Y in the database. Next we discuss the types of problems we may encounter with these two simple transactions if they run concurrently.

The Lost Update Problem. This problem occurs when two transactions that access the same database items have their operations interleaved in a way that makes the value of some database items incorrect. Suppose that transactions T_1 and T_2 are submitted at approximately the same time, and suppose that their operations are interleaved as shown in Figure 21.3(a); then the final value of item X is incorrect because T_2 reads the value of X *before* T_1 changes it in the database, and hence the updated value resulting from T_1 is lost. For example, if $X = 80$ at the start (originally there were 80 reservations on the flight), $N = 5$ (T_1 transfers 5 seat reservations from the flight corresponding to X to the flight corresponding to Y), and $M = 4$ (T_2 reserves 4 seats on X), the final result should be $X = 79$. However, in the interleaving of operations shown in Figure 21.3(a), it is $X = 84$ because the update in T_1 that removed the five seats from X was *lost*.

The Temporary Update (or Dirty Read) Problem. This problem occurs when one transaction updates a database item and then the transaction fails for some reason (see Section 21.1.4). Meanwhile, the updated item is accessed (read) by another transaction before it is changed back to its original value. Figure 21.3(b) shows an example where T_1 updates item X and then fails before completion, so the system must change X back to its original value. Before it can do so, however, transaction T_2 reads the *temporary* value of X, which will not be recorded permanently in the database because of the failure of T_1. The value of item X that is read by T_2 is called *dirty data* because it has been created by a transaction that has not completed and committed yet; hence, this problem is also known as the *dirty read problem*.

The Incorrect Summary Problem. If one transaction is calculating an aggregate summary function on a number of database items while other transactions are updating some of these items, the aggregate function may calculate some values before they are updated and others after they are updated. For example, suppose that a transaction T_3 is calculating the total number of reservations on all the flights; meanwhile, transaction T_1 is executing. If the interleaving of operations shown in Figure 21.3(c) occurs, the result of T_3 will be off by an amount N because T_3 reads the value of X *after* N seats have been subtracted from it but reads the value of Y *before* those N seats have been added to it.

Figure 21.3
Some problems that occur when concurrent execution is uncontrolled. (a) The lost update problem. (b) The temporary update problem. (c) The incorrect summary problem.

(a)

Time

T_1	T_2
read_item(X); $X := X - N$;	
	read_item(X); $X := X + M$;
write_item(X); read_item(Y);	
	write_item(X);
$Y := Y + N$; write_item(Y);	

Item X has an incorrect value because its update by T_1 is *lost* (overwritten).

(b)

Time

T_1	T_2

read_item(X);
$X := X - N$;
write_item(X);	

	read_item(X);
	$X := X + M$;
	write_item(X);

read_item(Y);	

Transaction T_1 fails and must change the value of X back to its old value; meanwhile T_2 has read the *temporary* incorrect value of X.

(c)

T_1	T_3
	$sum := 0$; read_item(A); $sum := sum + A$; ⋮

read_item(X);
$X := X - N$;
write_item(X);	

	read_item(X);
	$sum := sum + X$;
	read_item(Y);
	$sum := sum + Y$;

read_item(Y);
$Y := Y + N$;
write_item(Y);	

T_3 reads X after N is subtracted and reads Y before N is added; a wrong summary is the result (off by N).

The Unrepeatable Read Problem. Another problem that may occur is called *unrepeatable read*, where a transaction T reads the same item twice and the item is changed by another transaction T' between the two reads. Hence, T receives *different values* for its two reads of the same item. This may occur, for example, if during an airline reservation transaction, a customer inquires about seat availability on several flights. When the customer decides on a particular flight, the transaction then reads the number of seats on that flight a second time before completing the reservation, and it may end up reading a different value for the item.

21.1.4 Why Recovery Is Needed

Whenever a transaction is submitted to a DBMS for execution, the system is responsible for making sure that either all the operations in the transaction are completed successfully and their effect is recorded permanently in the database, or that the transaction does not have any effect on the database or any other transactions. In the first case, the transaction is said to be **committed**, whereas in the second case, the transaction is **aborted**. The DBMS must not permit some operations of a transaction T to be applied to the database while other operations of T are not, because *the whole transaction* is a logical unit of database processing. If a transaction **fails** after executing some of its operations but before executing all of them, the operations already executed must be undone and have no lasting effect.

Types of Failures. Failures are generally classified as transaction, system, and media failures. There are several possible reasons for a transaction to fail in the middle of execution:

1. **A computer failure (system crash).** A hardware, software, or network error occurs in the computer system during transaction execution. Hardware crashes are usually media failures—for example, main memory failure.
2. **A transaction or system error.** Some operation in the transaction may cause it to fail, such as integer overflow or division by zero. Transaction failure may also occur because of erroneous parameter values or because of a logical programming error.[3] Additionally, the user may interrupt the transaction during its execution.
3. **Local errors or exception conditions detected by the transaction.** During transaction execution, certain conditions may occur that necessitate cancellation of the transaction. For example, data for the transaction may not be found. An exception condition,[4] such as insufficient account balance in a banking database, may cause a transaction, such as a fund withdrawal, to be canceled. This exception could be programmed in the transaction itself, and in such a case would not be considered as a transaction failure.

[3]In general, a transaction should be thoroughly tested to ensure that it does not have any bugs (logical programming errors).

[4]Exception conditions, if programmed correctly, do not constitute transaction failures.

4. **Concurrency control enforcement.** The concurrency control method (see Chapter 22) may decide to abort a transaction because it violates serializability (see Section 21.5), or it may abort one or more transactions to resolve a state of deadlock among several transactions (see Section 22.1.3). Transactions aborted because of serializability violations or deadlocks are typically restarted automatically at a later time.
5. **Disk failure.** Some disk blocks may lose their data because of a read or write malfunction or because of a disk read/write head crash. This may happen during a read or a write operation of the transaction.
6. **Physical problems and catastrophes.** This refers to an endless list of problems that includes power or air-conditioning failure, fire, theft, sabotage, overwriting disks or tapes by mistake, and mounting of a wrong tape by the operator.

Failures of types 1, 2, 3, and 4 are more common than those of types 5 or 6. Whenever a failure of type 1 through 4 occurs, the system must keep sufficient information to quickly recover from the failure. Disk failure or other catastrophic failures of type 5 or 6 do not happen frequently; if they do occur, recovery is a major task. We discuss recovery from failure in Chapter 23.

The concept of transaction is fundamental to many techniques for concurrency control and recovery from failures.

21.2 Transaction and System Concepts

In this section we discuss additional concepts relevant to transaction processing. Section 21.2.1 describes the various states a transaction can be in, and discusses other operations needed in transaction processing. Section 21.2.2 discusses the system log, which keeps information about transactions and data items that will be needed for recovery. Section 21.2.3 describes the concept of commit points of transactions, and why they are important in transaction processing.

21.2.1 Transaction States and Additional Operations

A transaction is an atomic unit of work that should either be completed in its entirety or not done at all. For recovery purposes, the system needs to keep track of when each transaction starts, terminates, and commits or aborts (see Section 21.2.3). Therefore, the recovery manager of the DBMS needs to keep track of the following operations:

- BEGIN_TRANSACTION. This marks the beginning of transaction execution.
- READ or WRITE. These specify read or write operations on the database items that are executed as part of a transaction.
- END_TRANSACTION. This specifies that READ and WRITE transaction operations have ended and marks the end of transaction execution. However, at this point it may be necessary to check whether the changes introduced by

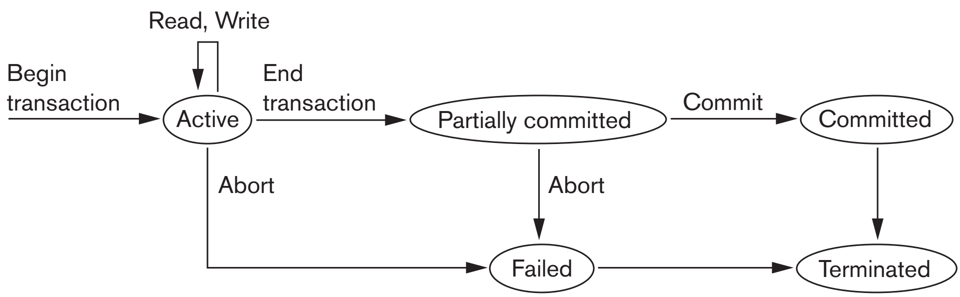
the transaction can be permanently applied to the database (committed) or whether the transaction has to be aborted because it violates serializability (see Section 21.5) or for some other reason.

- COMMIT_TRANSACTION. This signals a *successful end* of the transaction so that any changes (updates) executed by the transaction can be safely **committed** to the database and will not be undone.
- ROLLBACK (or ABORT). This signals that the transaction has *ended unsuccessfully,* so that any changes or effects that the transaction may have applied to the database must be **undone**.

Figure 21.4 shows a state transition diagram that illustrates how a transaction moves through its execution states. A transaction goes into an **active state** immediately after it starts execution, where it can execute its READ and WRITE operations. When the transaction ends, it moves to the **partially committed state**. At this point, some recovery protocols need to ensure that a system failure will not result in an inability to record the changes of the transaction permanently (usually by recording changes in the system log, discussed in the next section).[5] Once this check is successful, the transaction is said to have reached its commit point and enters the **committed state**. Commit points are discussed in more detail in Section 21.2.3. When a transaction is committed, it has concluded its execution successfully and all its changes must be recorded permanently in the database, even if a system failure occurs.

However, a transaction can go to the **failed state** if one of the checks fails or if the transaction is aborted during its active state. The transaction may then have to be rolled back to undo the effect of its WRITE operations on the database. The **terminated state** corresponds to the transaction leaving the system. The transaction information that is maintained in system tables while the transaction has been running is removed when the transaction terminates. Failed or aborted transactions may be *restarted* later—either automatically or after being resubmitted by the user—as brand new transactions.

Figure 21.4
State transition diagram illustrating the states for transaction execution.



[5]Optimistic concurrency control (see Section 22.4) also requires that certain checks are made at this point to ensure that the transaction did not interfere with other executing transactions.

21.2.2 The System Log

To be able to recover from failures that affect transactions, the system maintains a **log**[6] to keep track of all transaction operations that affect the values of database items, as well as other transaction information that may be needed to permit recovery from failures. The log is a sequential, append-only file that is kept on disk, so it is not affected by any type of failure except for disk or catastrophic failure. Typically, one (or more) main memory buffers hold the last part of the log file, so that log entries are first added to the main memory buffer. When the **log buffer** is filled, or when certain other conditions occur, the log buffer is *appended to the end of the log file on disk*. In addition, the log file from disk is periodically backed up to archival storage (tape) to guard against catastrophic failures. The following are the types of entries—called **log records**—that are written to the log file and the corresponding action for each log record. In these entries, *T* refers to a unique **transaction-id** that is generated automatically by the system for each transaction and that is used to identify each transaction:

1. [**start_transaction, *T***]. Indicates that transaction *T* has started execution.
2. [**write_item, *T*, *X*, *old_value*, *new_value***]. Indicates that transaction *T* has changed the value of database item *X* from *old_value* to *new_value.*
3. [**read_item, *T*, *X***]. Indicates that transaction *T* has read the value of database item *X*.
4. [**commit, *T***]. Indicates that transaction *T* has completed successfully, and affirms that its effect can be committed (recorded permanently) to the database.
5. [**abort, *T***]. Indicates that transaction *T* has been aborted.

Protocols for recovery that avoid cascading rollbacks (see Section 21.4.2)—which include nearly all practical protocols—*do not require* that READ operations are written to the system log. However, if the log is also used for other purposes—such as auditing (keeping track of all database operations)—then such entries can be included. Additionally, some recovery protocols require simpler WRITE entries only include one of new_value and old_value instead of including both (see Section 21.4.2).

Notice that we are assuming that all permanent changes to the database occur within transactions, so the notion of recovery from a transaction failure amounts to either undoing or redoing transaction operations individually from the log. If the system crashes, we can recover to a consistent database state by examining the log and using one of the techniques described in Chapter 23. Because the log contains a record of every WRITE operation that changes the value of some database item, it is possible to **undo** the effect of these WRITE operations of a transaction *T* by tracing backward through the log and resetting all items changed by a WRITE operation of *T* to their old_values. **Redo** of an operation may also be necessary if a transaction has its updates recorded in the log but a failure occurs before the system can be sure that

[6]The log has sometimes been called the *DBMS journal*.

all these new_values have been written to the actual database on disk from the main memory buffers.[7]

21.2.3 Commit Point of a Transaction

A transaction *T* reaches its **commit point** when all its operations that access the database have been executed successfully *and* the effect of all the transaction operations on the database have been recorded in the log. Beyond the commit point, the transaction is said to be **committed**, and its effect must be *permanently recorded* in the database. The transaction then writes a commit record [commit, *T*] into the log. If a system failure occurs, we can search back in the log for all transactions *T* that have written a [start_transaction, *T*] record into the log but have not written their [commit, *T*] record yet; these transactions may have to be *rolled back* to *undo their effect* on the database during the recovery process. Transactions that have written their commit record in the log must also have recorded all their WRITE operations in the log, so their effect on the database can be *redone* from the log records.

Notice that the log file must be kept on disk. As discussed in Chapter 17, updating a disk file involves copying the appropriate block of the file from disk to a buffer in main memory, updating the buffer in main memory, and copying the buffer to disk. It is common to keep one or more blocks of the log file in main memory buffers, called the **log buffer**, until they are filled with log entries and then to write them back to disk only once, rather than writing to disk every time a log entry is added. This saves the overhead of multiple disk writes of the same log file buffer. At the time of a system crash, only the log entries that have been *written back to disk* are considered in the recovery process because the contents of main memory may be lost. Hence, *before* a transaction reaches its commit point, any portion of the log that has not been written to the disk yet must now be written to the disk. This process is called **force-writing** the log buffer before committing a transaction.

21.3 Desirable Properties of Transactions

Transactions should possess several properties, often called the **ACID** properties; they should be enforced by the concurrency control and recovery methods of the DBMS. The following are the ACID properties:

- **Atomicity.** A transaction is an atomic unit of processing; it should either be performed in its entirety or not performed at all.
- **Consistency preservation.** A transaction should be consistency preserving, meaning that if it is completely executed from beginning to end without interference from other transactions, it should take the database from one consistent state to another.
- **Isolation.** A transaction should appear as though it is being executed in isolation from other transactions, even though many transactions are executing

[7]Undo and redo are discussed more fully in Chapter 23.

concurrently. That is, the execution of a transaction should not be interfered with by any other transactions executing concurrently.

- **Durability or permanency.** The changes applied to the database by a committed transaction must persist in the database. These changes must not be lost because of any failure.

The *atomicity property* requires that we execute a transaction to completion. It is the responsibility of the *transaction recovery subsystem* of a DBMS to ensure atomicity. If a transaction fails to complete for some reason, such as a system crash in the midst of transaction execution, the recovery technique must undo any effects of the transaction on the database. On the other hand, write operations of a committed transaction must be eventually written to disk.

The preservation of *consistency* is generally considered to be the responsibility of the programmers who write the database programs or of the DBMS module that enforces integrity constraints. Recall that a **database state** is a collection of all the stored data items (values) in the database at a given point in time. A **consistent state** of the database satisfies the constraints specified in the schema as well as any other constraints on the database that should hold. A database program should be written in a way that guarantees that, if the database is in a consistent state before executing the transaction, it will be in a consistent state after the *complete* execution of the transaction, assuming that *no interference with other transactions* occurs.

The *isolation property* is enforced by the *concurrency control subsystem* of the DBMS.[8] If every transaction does not make its updates (write operations) visible to other transactions until it is committed, one form of isolation is enforced that solves the temporary update problem and eliminates cascading rollbacks (see Chapter 23) but does not eliminate all other problems. There have been attempts to define the **level of isolation** of a transaction. A transaction is said to have level 0 (zero) isolation if it does not overwrite the dirty reads of higher-level transactions. Level 1 (one) isolation has no lost updates, and level 2 isolation has no lost updates and no dirty reads. Finally, level 3 isolation (also called *true isolation*) has, in addition to level 2 properties, repeatable reads.[9]

And last, the *durability property* is the responsibility of the *recovery subsystem* of the DBMS. We will introduce how recovery protocols enforce durability and atomicity in the next section and then discuss this in more detail in Chapter 23.

21.4 Characterizing Schedules Based on Recoverability

When transactions are executing concurrently in an interleaved fashion, then the order of execution of operations from all the various transactions is known as a **schedule** (or **history**). In this section, first we define the concept of schedules, and

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[8]We will discuss concurrency control protocols in Chapter 22.

[9]The SQL syntax for isolation level discussed later in Section 21.6 is closely related to these levels.

then we characterize the types of schedules that facilitate recovery when failures occur. In Section 21.5, we characterize schedules in terms of the interference of participating transactions, leading to the concepts of serializability and serializable schedules.

21.4.1 Schedules (Histories) of Transactions

A **schedule** (or **history**) S of n transactions $T_1, T_2, ..., T_n$ is an ordering of the operations of the transactions. Operations from different transactions can be interleaved in the schedule S. However, for each transaction T_i that participates in the schedule S, the operations of T_i in S must appear in the same order in which they occur in T_i. The order of operations in S is considered to be a *total ordering*, meaning *that for any two operations* in the schedule, one must occur before the other. It is possible theoretically to deal with schedules whose operations form *partial orders* (as we discuss later), but we will assume for now total ordering of the operations in a schedule.

For the purpose of recovery and concurrency control, we are mainly interested in the read_item and write_item operations of the transactions, as well as the commit and abort operations. A shorthand notation for describing a schedule uses the symbols b, r, w, e, c, and a for the operations begin_transaction, read_item, write_item, end_transaction, commit, and abort, respectively, and appends as a *subscript* the transaction id (transaction number) to each operation in the schedule. In this notation, the database item X that is read or written follows the r and w operations in parentheses. In some schedules, we will only show the *read* and *write* operations, whereas in other schedules, we will show all the operations. For example, the schedule in Figure 21.3(a), which we shall call S_a, can be written as follows in this notation:

$$S_a: r_1(X); r_2(X); w_1(X); r_1(Y); w_2(X); w_1(Y);$$

Similarly, the schedule for Figure 21.3(b), which we call S_b, can be written as follows, if we assume that transaction T_1 aborted after its read_item(Y) operation:

$$S_b: r_1(X); w_1(X); r_2(X); w_2(X); r_1(Y); a_1;$$

Two operations in a schedule are said to **conflict** if they satisfy all three of the following conditions: (1) they belong to *different transactions*; (2) they access the *same item X*; and (3) *at least one* of the operations is a write_item(X). For example, in schedule S_a, the operations $r_1(X)$ and $w_2(X)$ conflict, as do the operations $r_2(X)$ and $w_1(X)$, and the operations $w_1(X)$ and $w_2(X)$. However, the operations $r_1(X)$ and $r_2(X)$ do not conflict, since they are both read operations; the operations $w_2(X)$ and $w_1(Y)$ do not conflict because they operate on distinct data items X and Y; and the operations $r_1(X)$ and $w_1(X)$ do not conflict because they belong to the same transaction.

Intuitively, two operations are conflicting if changing their order can result in a different outcome. For example, if we change the order of the two operations $r_1(X)$; $w_2(X)$ to $w_2(X)$; $r_1(X)$, then the value of X that is read by transaction T_1 changes, because in the second order the value of X is changed by $w_2(X)$ before it is read by

$r_1(X)$, whereas in the first order the value is read before it is changed. This is called a **read-write conflict**. The other type is called a **write-write conflict**, and is illustrated by the case where we change the order of two operations such as $w_1(X)$; $w_2(X)$ to $w_2(X)$; $w_1(X)$. For a write-write conflict, the *last value* of X will differ because in one case it is written by T_2 and in the other case by T_1. Notice that two read operations are not conflicting because changing their order makes no difference in outcome.

The rest of this section covers some theoretical definitions concerning schedules. A schedule S of n transactions $T_1, T_2, ..., T_n$ is said to be a **complete schedule** if the following conditions hold:

1. The operations in S are exactly those operations in $T_1, T_2, ..., T_n$, including a commit or abort operation as the last operation for each transaction in the schedule.
2. For any pair of operations from the same transaction T_i, their relative order of appearance in S is the same as their order of appearance in T_i.
3. For any two conflicting operations, one of the two must occur before the other in the schedule.[10]

The preceding condition (3) allows for two *nonconflicting operations* to occur in the schedule without defining which occurs first, thus leading to the definition of a schedule as a **partial order** of the operations in the n transactions.[11] However, a total order must be specified in the schedule for any pair of conflicting operations (condition 3) and for any pair of operations from the same transaction (condition 2). Condition 1 simply states that all operations in the transactions must appear in the complete schedule. Since every transaction has either committed or aborted, a complete schedule will *not contain any active transactions* at the end of the schedule.

In general, it is difficult to encounter complete schedules in a transaction processing system because new transactions are continually being submitted to the system. Hence, it is useful to define the concept of the **committed projection** $C(S)$ of a schedule S, which includes only the operations in S that belong to committed transactions—that is, transactions T_i whose commit operation c_i is in S.

21.4.2 Characterizing Schedules Based on Recoverability

For some schedules it is easy to recover from transaction and system failures, whereas for other schedules the recovery process can be quite involved. In some cases, it is even not possible to recover correctly after a failure. Hence, it is important to characterize the types of schedules for which *recovery is possible*, as well as those for which *recovery is relatively simple*. These characterizations do not actually provide the recovery algorithm; they only attempt to theoretically characterize the different types of schedules.

[10]Theoretically, it is not necessary to determine an order between pairs of *nonconflicting* operations.

[11]In practice, most schedules have a total order of operations. If parallel processing is employed, it is theoretically possible to have schedules with partially ordered nonconflicting operations.

First, we would like to ensure that, once a transaction T is committed, it should *never* be necessary to roll back T. This ensures that the durability property of transactions is not violated (see Section 21.3). The schedules that theoretically meet this criterion are called *recoverable schedules;* those that do not are called **nonrecoverable** and hence should not be permitted by the DBMS. The definition of **recoverable schedule** is as follows: A schedule S is recoverable if no transaction T in S commits until all transactions T' that have written some item X that T reads have committed. A transaction T **reads** from transaction T' in a schedule S if some item X is first written by T' and later read by T. In addition, T' should not have been aborted before T reads item X, and there should be no transactions that write X after T' writes it and before T reads it (unless those transactions, if any, have aborted before T reads X).

Some recoverable schedules may require a complex recovery process as we shall see, but if sufficient information is kept (in the log), a recovery algorithm can be devised for any recoverable schedule. The (partial) schedules S_a and S_b from the preceding section are both recoverable, since they satisfy the above definition. Consider the schedule S_a' given below, which is the same as schedule S_a except that two commit operations have been added to S_a:

S_a': $r_1(X); r_2(X); w_1(X); r_1(Y); w_2(X); c_2; w_1(Y); c_1;$

S_a' is recoverable, even though it suffers from the lost update problem; this problem is handled by serializability theory (see Section 21.5). However, consider the two (partial) schedules S_c and S_d that follow:

S_c: $r_1(X); w_1(X); r_2(X); r_1(Y); w_2(X); c_2; a_1;$
S_d: $r_1(X); w_1(X); r_2(X); r_1(Y); w_2(X); w_1(Y); c_1; c_2;$
S_e: $r_1(X); w_1(X); r_2(X); r_1(Y); w_2(X); w_1(Y); a_1; a_2;$

S_c is not recoverable because T_2 reads item X from T_1, but T_2 commits before T_1 commits. The problem occurs if T_1 aborts after the c_2 operation in S_c, then the value of X that T_2 read is no longer valid and T_2 must be aborted *after* it is committed, leading to a schedule that is *not recoverable*. For the schedule to be recoverable, the c_2 operation in S_c must be postponed until after T_1 commits, as shown in S_d. If T_1 aborts instead of committing, then T_2 should also abort as shown in S_e, because the value of X it read is no longer valid. In S_e, aborting T_2 is acceptable since it has not committed yet, which is not the case for the nonrecoverable schedule S_c.

In a recoverable schedule, no committed transaction ever needs to be rolled back, and so the definition of committed transaction as durable is not violated. However, it is possible for a phenomenon known as **cascading rollback** (or **cascading abort**) to occur in some recoverable schedules, where an *uncommitted* transaction has to be rolled back because it read an item from a transaction that failed. This is illustrated in schedule S_e, where transaction T_2 has to be rolled back because it read item X from T_1, and T_1 then aborted.

Because cascading rollback can be quite time-consuming—since numerous transactions can be rolled back (see Chapter 23)—it is important to characterize the schedules where this phenomenon is guaranteed not to occur. A schedule is said to be **cascadeless**, or to **avoid cascading rollback**, if every transaction in the schedule reads only items that were written by committed transactions. In this case, all items read will not be discarded, so no cascading rollback will occur. To satisfy this criterion, the $r_2(X)$ command in schedules S_d and S_e must be postponed until after T_1 has committed (or aborted), thus delaying T_2 but ensuring no cascading rollback if T_1 aborts.

Finally, there is a third, more restrictive type of schedule, called a **strict schedule**, in which transactions can *neither read nor write* an item X until the last transaction that wrote X has committed (or aborted). Strict schedules simplify the recovery process. In a strict schedule, the process of undoing a write_item(X) operation of an aborted transaction is simply to restore the **before image** (old_value or BFIM) of data item X. This simple procedure always works correctly for strict schedules, but it may not work for recoverable or cascadeless schedules. For example, consider schedule S_f:

S_f: $w_1(X, 5)$; $w_2(X, 8)$; a_1;

Suppose that the value of X was originally 9, which is the before image stored in the system log along with the $w_1(X, 5)$ operation. If T_1 aborts, as in S_f, the recovery procedure that restores the before image of an aborted write operation will restore the value of X to 9, even though it has already been changed to 8 by transaction T_2, thus leading to potentially incorrect results. Although schedule S_f is cascadeless, it is not a strict schedule, since it permits T_2 to write item X even though the transaction T_1 that last wrote X had not yet committed (or aborted). A strict schedule does not have this problem.

It is important to note that any strict schedule is also cascadeless, and any cascadeless schedule is also recoverable. Suppose we have i transactions $T_1, T_2, \ldots, T_i$, and their number of operations are $n_1, n_2, \ldots, n_i$, respectively. If we make a set of all possible schedules of these transactions, we can divide the schedules into two disjoint subsets: recoverable and nonrecoverable. The cascadeless schedules will be a subset of the recoverable schedules, and the strict schedules will be a subset of the cascadeless schedules. Thus, all strict schedules are cascadeless, and all cascadeless schedules are recoverable.

21.5 Characterizing Schedules Based on Serializability

In the previous section, we characterized schedules based on their recoverability properties. Now we characterize the types of schedules that are always considered to be *correct* when concurrent transactions are executing. Such schedules are known as *serializable schedules*. Suppose that two users—for example, two airline reservations agents—submit to the DBMS transactions T_1 and T_2 in Figure 21.2 at approximately the same time. If no interleaving of operations is permitted, there are only two possible outcomes:

1. Execute all the operations of transaction T_1 (in sequence) followed by all the operations of transaction T_2 (in sequence).

2. Execute all the operations of transaction T_2 (in sequence) followed by all the operations of transaction T_1 (in sequence).

These two schedules—called *serial schedules*—are shown in Figure 21.5(a) and (b), respectively. If interleaving of operations is allowed, there will be many possible orders in which the system can execute the individual operations of the transactions. Two possible schedules are shown in Figure 21.5(c). The concept of **serializability of schedules** is used to identify which schedules are correct when transaction executions have interleaving of their operations in the schedules. This section defines serializability and discusses how it may be used in practice.

Figure 21.5
Examples of serial and nonserial schedules involving transactions T_1 and T_2. (a) Serial schedule A: T_1 followed by T_2. (b) Serial schedule B: T_2 followed by T_1. (c) Two nonserial schedules C and D with interleaving of operations.

(a)

Time

T_1	T_2
read_item(X);	
$X := X - N$;	
write_item(X);	
read_item(Y);	
$Y := Y + N$;	
write_item(Y);	
	read_item(X);
	$X := X + M$;
	write_item(X);

Schedule A

(b)

Time

T_1	T_2
	read_item(X);
	$X := X + M$;
	write_item(X);
read_item(X);	
$X := X - N$;	
write_item(X);	
read_item(Y);	
$Y := Y + N$;	
write_item(Y);	

Schedule B

(c)

Time

T_1	T_2
read_item(X);	
$X := X - N$;	
	read_item(X);
	$X := X + M$;
write_item(X);	
read_item(Y);	
	write_item(X);
$Y := Y + N$;	
write_item(Y);	

Schedule C

Time

T_1	T_2
read_item(X);	
$X := X - N$;	
write_item(X);	
	read_item(X);
	$X := X + M$;
	write_item(X);
read_item(Y);	
$Y := Y + N$;	
write_item(Y);	

Schedule D

21.5.1 Serial, Nonserial, and Conflict-Serializable Schedules

Schedules A and B in Figure 21.5(a) and (b) are called *serial* because the operations of each transaction are executed consecutively, without any interleaved operations from the other transaction. In a serial schedule, entire transactions are performed in serial order: T_1 and then T_2 in Figure 21.5(a), and T_2 and then T_1 in Figure 21.5(b). Schedules C and D in Figure 21.5(c) are called *nonserial* because each sequence interleaves operations from the two transactions.

Formally, a schedule S is **serial** if, for every transaction T participating in the schedule, all the operations of T are executed consecutively in the schedule; otherwise, the schedule is called **nonserial**. Therefore, in a serial schedule, only one transaction at a time is active—the commit (or abort) of the active transaction initiates execution of the next transaction. No interleaving occurs in a serial schedule. One reasonable assumption we can make, if we consider the transactions to be *independent*, is that *every serial schedule is considered correct*. We can assume this because every transaction is assumed to be correct if executed on its own (according to the *consistency preservation* property of Section 21.3). Hence, it does not matter which transaction is executed first. As long as every transaction is executed from beginning to end in isolation from the operations of other transactions, we get a correct end result on the database.

The problem with serial schedules is that they limit concurrency by prohibiting interleaving of operations. In a serial schedule, if a transaction waits for an I/O operation to complete, we cannot switch the CPU processor to another transaction, thus wasting valuable CPU processing time. Additionally, if some transaction T is quite long, the other transactions must wait for T to complete all its operations before starting. Hence, serial schedules are *considered unacceptable* in practice. However, if we can determine which other schedules are *equivalent* to a serial schedule, we can allow these schedules to occur.

To illustrate our discussion, consider the schedules in Figure 21.5, and assume that the initial values of database items are $X = 90$ and $Y = 90$ and that $N = 3$ and $M = 2$. After executing transactions T_1 and T_2, we would expect the database values to be $X = 89$ and $Y = 93$, according to the meaning of the transactions. Sure enough, executing either of the serial schedules A or B gives the correct results. Now consider the nonserial schedules C and D. Schedule C (which is the same as Figure 21.3(a)) gives the results $X = 92$ and $Y = 93$, in which the X value is erroneous, whereas schedule D gives the correct results.

Schedule C gives an erroneous result because of the *lost update problem* discussed in Section 21.1.3; transaction T_2 reads the value of X before it is changed by transaction T_1, so only the effect of T_2 on X is reflected in the database. The effect of T_1 on X is *lost*, overwritten by T_2, leading to the incorrect result for item X. However, some nonserial schedules give the correct expected result, such as schedule D. We would like to determine which of the nonserial schedules *always* give a correct result and which may give erroneous results. The concept used to characterize schedules in this manner is that of serializability of a schedule.

The definition of *serializable schedule* is as follows: A schedule S of n transactions is **serializable** if it is *equivalent to some serial schedule* of the same n transactions. We will define the concept of *equivalence of schedules* shortly. Notice that there are $n!$ possible serial schedules of n transactions and many more possible nonserial schedules. We can form two disjoint groups of the nonserial schedules—those that are equivalent to one (or more) of the serial schedules and hence are serializable, and those that are not equivalent to *any* serial schedule and hence are not serializable.

Saying that a nonserial schedule S is serializable is equivalent to saying that it is correct, because it is equivalent to a serial schedule, which is considered correct. The remaining question is: When are two schedules considered *equivalent*?

There are several ways to define schedule equivalence. The simplest but least satisfactory definition involves comparing the effects of the schedules on the database. Two schedules are called **result equivalent** if they produce the same final state of the database. However, two different schedules may accidentally produce the same final state. For example, in Figure 21.6, schedules S_1 and S_2 will produce the same final database state if they execute on a database with an initial value of $X = 100$; however, for other initial values of X, the schedules are *not* result equivalent. Additionally, these schedules execute different transactions, so they definitely should not be considered equivalent. Hence, result equivalence alone cannot be used to define equivalence of schedules. The safest and most general approach to defining schedule equivalence is not to make any assumptions about the types of operations included in the transactions. For two schedules to be equivalent, the operations applied to each data item affected by the schedules should be applied to that item in both schedules *in the same order*. Two definitions of equivalence of schedules are generally used: *conflict equivalence* and *view equivalence*. We discuss conflict equivalence next, which is the more commonly used definition.

The definition of *conflict equivalence* of schedules is as follows: Two schedules are said to be **conflict equivalent** if the order of any two *conflicting operations* is the same in both schedules. Recall from Section 21.4.1 that two operations in a schedule are said to *conflict* if they belong to different transactions, access the same database item, and either both are write_item operations or one is a write_item and the other a read_item. If two conflicting operations are applied in *different orders* in two schedules, the effect can be different on the database or on the transactions in the schedule, and hence the schedules are not conflict equivalent. For example, as we discussed in Section 21.4.1, if a read and write operation occur in the order $r_1(X)$, $w_2(X)$ in schedule S_1, and in the reverse order $w_2(X)$, $r_1(X)$ in schedule S_2, the value read by $r_1(X)$ can be different in the two schedules. Similarly, if two write operations

Figure 21.6
Two schedules that are result equivalent for the initial value of $X = 100$ but are not result equivalent in general.

S_1

read_item(X);
$X := X + 10$;
write_item(X);

S_2
read_item(X); $X := X * 1.1$; write_item (X);

occur in the order $w_1(X)$, $w_2(X)$ in S_1, and in the reverse order $w_2(X)$, $w_1(X)$ in S_2, the next $r(X)$ operation in the two schedules will read potentially different values; or if these are the last operations writing item X in the schedules, the final value of item X in the database will be different.

Using the notion of conflict equivalence, we define a schedule S to be **conflict serializable**[12] if it is (conflict) equivalent to some serial schedule S'. In such a case, we can reorder the *nonconflicting* operations in S until we form the equivalent serial schedule S'. According to this definition, schedule D in Figure 21.5(c) is equivalent to the serial schedule A in Figure 21.5(a). In both schedules, the read_item(X) of T_2 reads the value of X written by T_1, while the other read_item operations read the database values from the initial database state. Additionally, T_1 is the last transaction to write Y, and T_2 is the last transaction to write X in both schedules. Because A is a serial schedule and schedule D is equivalent to A, D is a serializable schedule. Notice that the operations $r_1(Y)$ and $w_1(Y)$ of schedule D do not conflict with the operations $r_2(X)$ and $w_2(X)$, since they access different data items. Therefore, we can move $r_1(Y)$, $w_1(Y)$ before $r_2(X)$, $w_2(X)$, leading to the equivalent serial schedule T_1, T_2.

Schedule C in Figure 21.5(c) is not equivalent to either of the two possible serial schedules A and B, and hence is *not serializable*. Trying to reorder the operations of schedule C to find an equivalent serial schedule fails because $r_2(X)$ and $w_1(X)$ conflict, which means that we cannot move $r_2(X)$ down to get the equivalent serial schedule T_1, T_2. Similarly, because $w_1(X)$ and $w_2(X)$ conflict, we cannot move $w_1(X)$ down to get the equivalent serial schedule T_2, T_1.

Another, more complex definition of equivalence—called *view equivalence*, which leads to the concept of view serializability—is discussed in Section 21.5.4.

21.5.2 Testing for Conflict Serializability of a Schedule

There is a simple algorithm for determining whether a particular schedule is conflict serializable or not. Most concurrency control methods do *not* actually test for serializability. Rather protocols, or rules, are developed that guarantee that any schedule that follows these rules will be serializable. We discuss the algorithm for testing conflict serializability of schedules here to gain a better understanding of these concurrency control protocols, which are discussed in Chapter 22.

Algorithm 21.1 can be used to test a schedule for conflict serializability. The algorithm looks at only the read_item and write_item operations in a schedule to construct a **precedence graph** (or **serialization graph**), which is a **directed graph** $G = (N, E)$ that consists of a set of nodes $N = \{T_1, T_2, ..., T_n\}$ and a set of directed edges $E = \{e_1, e_2, ..., e_m\}$. There is one node in the graph for each transaction T_i in the schedule. Each edge e_i in the graph is of the form $(T_j \rightarrow T_k)$, $1 \leq j \leq n$, $1 \leq k \leq n$, where T_j is the **starting node** of e_i and T_k is the **ending node** of e_i. Such an edge from node T_j to

[12]We will use *serializable* to mean conflict serializable. Another definition of serializable used in practice (see Section 21.6) is to have repeatable reads, no dirty reads, and no phantom records (see Section 22.7.1 for a discussion on phantoms).

node T_k is created by the algorithm if one of the operations in T_j appears in the schedule before some *conflicting operation* in T_k.

Algorithm 21.1. Testing Conflict Serializability of a Schedule S

1. For each transaction T_i participating in schedule S, create a node labeled T_i in the precedence graph.
2. For each case in S where T_j executes a read_item(X) after T_i executes a write_item(X), create an edge ($T_i \rightarrow T_j$) in the precedence graph.
3. For each case in S where T_j executes a write_item(X) after T_i executes a read_item(X), create an edge ($T_i \rightarrow T_j$) in the precedence graph.
4. For each case in S where T_j executes a write_item(X) after T_i executes a write_item(X), create an edge ($T_i \rightarrow T_j$) in the precedence graph.
5. The schedule S is serializable if and only if the precedence graph has no cycles.

The precedence graph is constructed as described in Algorithm 21.1. If there is a cycle in the precedence graph, schedule S is not (conflict) serializable; if there is no cycle, S is serializable. A **cycle** in a directed graph is a **sequence of edges** $C = ((T_j \rightarrow T_k), (T_k \rightarrow T_p), \ldots, (T_i \rightarrow T_j))$ with the property that the starting node of each edge—except the first edge—is the same as the ending node of the previous edge, and the starting node of the first edge is the same as the ending node of the last edge (the sequence starts and ends at the same node).

In the precedence graph, an edge from T_i to T_j means that transaction T_i must come before transaction T_j in any serial schedule that is equivalent to S, because two conflicting operations appear in the schedule in that order. If there is no cycle in the precedence graph, we can create an **equivalent serial schedule** S' that is equivalent to S, by ordering the transactions that participate in S as follows: Whenever an edge exists in the precedence graph from T_i to T_j, T_i must appear before T_j in the equivalent serial schedule S'.[13] Notice that the edges ($T_i \rightarrow T_j$) in a precedence graph can optionally be labeled by the name(s) of the data item(s) that led to creating the edge. Figure 21.7 shows such labels on the edges.

In general, several serial schedules can be equivalent to S if the precedence graph for S has no cycle. However, if the precedence graph has a cycle, it is easy to show that we cannot create any equivalent serial schedule, so S is not serializable. The precedence graphs created for schedules A to D, respectively, in Figure 21.5 appear in Figure 21.7(a) to (d). The graph for schedule C has a cycle, so it is not serializable. The graph for schedule D has no cycle, so it is serializable, and the equivalent serial schedule is T_1 followed by T_2. The graphs for schedules A and B have no cycles, as expected, because the schedules are serial and hence serializable.

Another example, in which three transactions participate, is shown in Figure 21.8. Figure 21.8(a) shows the read_item and write_item operations in each transaction. Two schedules E and F for these transactions are shown in Figure 21.8(b) and (c),

[13]This process of ordering the nodes of an acrylic graph is known as *topological sorting*.

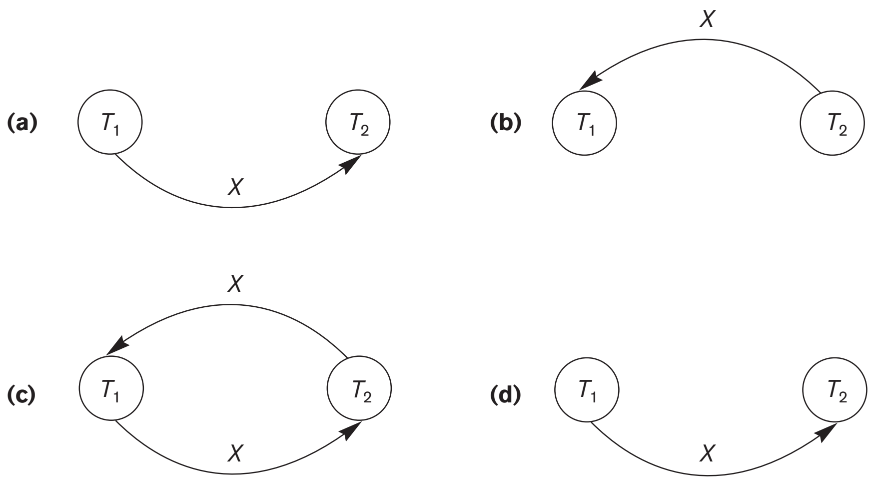


Figure 21.7
Constructing the precedence graphs for schedules A to D from Figure 21.5 to test for conflict serializability. (a) Precedence graph for serial schedule A. (b) Precedence graph for serial schedule B. (c) Precedence graph for schedule C (not serializable). (d) Precedence graph for schedule D (serializable, equivalent to schedule A).

respectively, and the precedence graphs for schedules E and F are shown in parts (d) and (e). Schedule E is not serializable because the corresponding precedence graph has cycles. Schedule F is serializable, and the serial schedule equivalent to *F* is shown in Figure 21.8(e). Although only one equivalent serial schedule exists for *F*, in general there may be more than one equivalent serial schedule for a serializable schedule. Figure 21.8(f) shows a precedence graph representing a schedule that has two equivalent serial schedules. To find an equivalent serial schedule, start with a node that does not have any incoming edges, and then make sure that the node order for every edge is not violated.

21.5.3 How Serializability Is Used for Concurrency Control

As we discussed earlier, saying that a schedule *S* is (conflict) serializable—that is, *S* is (conflict) equivalent to a serial schedule—is tantamount to saying that *S* is correct. Being *serializable* is distinct from being *serial*, however. A serial schedule represents inefficient processing because no interleaving of operations from different transactions is permitted. This can lead to low CPU utilization while a transaction waits for disk I/O, or for another transaction to terminate, thus slowing down processing considerably. A serializable schedule gives the benefits of concurrent execution without giving up any correctness. In practice, it is quite difficult to test for the serializability of a schedule. The interleaving of operations from concurrent transactions—which are usually executed as processes by the operating system—is typically determined by the operating system scheduler, which allocates resources to

(a)

Transaction T_1	Transaction T_2	Transaction T_3
read_item(X);	read_item(Z);	read_item(Y);
write_item(X);	read_item(Y);	read_item(Z);
read_item(Y);	write_item(Y);	write_item(Y);
write_item(Y);	read_item(X);	write_item(Z);
	write_item(X);	

(b)

Time ↓

Transaction T_1	Transaction T_2	Transaction T_3

	read_item(Z);
	read_item(Y);
	write_item(Y);	

		read_item(Y); read_item(Z);
read_item(X); write_item(X);		
		write_item(Y); write_item(Z);
	read_item(X);	
read_item(Y); write_item(Y);		
	write_item(X);	

Schedule E

(c)

Time ↓

Transaction T_1	Transaction T_2	Transaction T_3
		read_item(Y); read_item(Z);
read_item(X); write_item(X);		
		write_item(Y); write_item(Z);
	read_item(Z);	
read_item(Y); write_item(Y);		

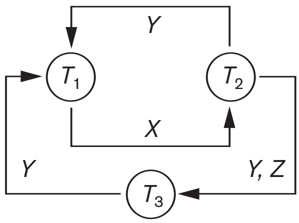
	read_item(Y);
	write_item(Y);
	read_item(X);
	write_item(X);	

Schedule F

Figure 21.8
Another example of serializability testing. (a) The read and write operations of three transactions T_1, T_2, and T_3. (b) Schedule E. (c) Schedule F.

all processes. Factors such as system load, time of transaction submission, and priorities of processes contribute to the ordering of operations in a schedule. Hence, it is difficult to determine how the operations of a schedule will be interleaved beforehand to ensure serializability.

(d)



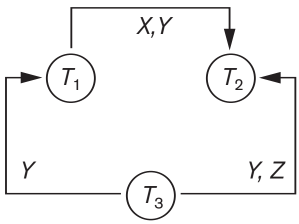
Equivalent serial schedules

None

Reason

Cycle $X(T_1 \rightarrow T_2), Y(T_2 \rightarrow T_1)$
Cycle $X(T_1 \rightarrow T_2), YZ\ (T_2 \rightarrow T_3), Y(T_3 \rightarrow T_1)$

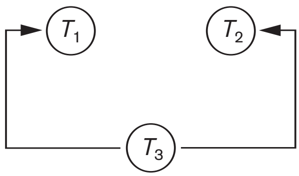
(e)



Equivalent serial schedules

$T_3 \rightarrow T_1 \rightarrow T_2$

(f)



Equivalent serial schedules

$T_3 \rightarrow T_1 \rightarrow T_2$

$T_3 \rightarrow T_2 \rightarrow T_1$

Figure 21.8 (continued)
Another example of serializability testing.
(d) Precedence graph for schedule E.
(e) Precedence graph for schedule F.
(f) Precedence graph with two equivalent serial schedules.

If transactions are executed at will and then the resulting schedule is tested for serializability, we must cancel the effect of the schedule if it turns out not to be serializable. This is a serious problem that makes this approach impractical. Hence, the approach taken in most practical systems is to determine methods or protocols that ensure serializability, without having to test the schedules themselves. The approach taken in most commercial DBMSs is to design **protocols** (sets of rules) that—if followed by *every* individual transaction or if enforced by a DBMS concurrency control subsystem—will ensure serializability of *all schedules in which the transactions participate*.

Another problem appears here: When transactions are submitted continuously to the system, it is difficult to determine when a schedule begins and when it ends. Serializability theory can be adapted to deal with this problem by considering only the committed projection of a schedule *S*. Recall from Section 21.4.1 that the *committed projection* $C(S)$ of a schedule *S* includes only the operations in *S* that belong to committed transactions. We can theoretically define a schedule *S* to be serializable if its committed projection $C(S)$ is equivalent to some serial schedule, since only committed transactions are guaranteed by the DBMS.

In Chapter 22, we discuss a number of different concurrency control protocols that guarantee serializability. The most common technique, called *two-phase locking*, is based on locking data items to prevent concurrent transactions from interfering with one another, and enforcing an additional condition that guarantees serializability. This is used in the majority of commercial DBMSs. Other protocols have been proposed;[14] these include *timestamp ordering*, where each transaction is assigned a unique timestamp and the protocol ensures that any conflicting operations are executed in the order of the transaction timestamps; *multiversion protocols*, which are based on maintaining multiple versions of data items; and *optimistic* (also called *certification* or *validation*) *protocols*, which check for possible serializability violations after the transactions terminate but before they are permitted to commit.

21.5.4 View Equivalence and View Serializability

In Section 21.5.1 we defined the concepts of conflict equivalence of schedules and conflict serializability. Another less restrictive definition of equivalence of schedules is called *view equivalence*. This leads to another definition of serializability called *view serializability*. Two schedules S and S' are said to be **view equivalent** if the following three conditions hold:

1. The same set of transactions participates in S and S', and S and S' include the same operations of those transactions.
2. For any operation $r_i(X)$ of T_i in S, if the value of X read by the operation has been written by an operation $w_j(X)$ of T_j (or if it is the original value of X before the schedule started), the same condition must hold for the value of X read by operation $r_i(X)$ of T_i in S'.
3. If the operation $w_k(Y)$ of T_k is the last operation to write item Y in S, then $w_k(Y)$ of T_k must also be the last operation to write item Y in S'.

The idea behind view equivalence is that, as long as each read operation of a transaction reads the result of the same write operation in both schedules, the write operations of each transaction must produce the same results. The read operations are hence said to *see the same view* in both schedules. Condition 3 ensures that the final write operation on each data item is the same in both schedules, so the database state should be the same at the end of both schedules. A schedule S is said to be **view serializable** if it is view equivalent to a serial schedule.

The definitions of conflict serializability and view serializability are similar if a condition known as the **constrained write assumption** (or **no blind writes**) holds on all transactions in the schedule. This condition states that any write operation $w_i(X)$ in T_i is preceded by a $r_i(X)$ in T_i and that the value written by $w_i(X)$ in T_i depends only on the value of X read by $r_i(X)$. This assumes that computation of the new value of X is a function $f(X)$ based on the old value of X read from the database. A **blind write** is a write operation in a transaction T on an item X that is not dependent on the value of X, so it is not preceded by a read of X in the transaction T.

[14]These other protocols have not been incorporated much into commercial systems; most relational DBMSs use some variation of the two-phase locking protocol.

The definition of view serializability is less restrictive than that of conflict serializability under the **unconstrained write assumption**, where the value written by an operation $w_i(X)$ in T_i can be independent of its old value from the database. This is possible when *blind writes* are allowed, and it is illustrated by the following schedule S_g of three transactions T_1: $r_1(X)$; $w_1(X)$; T_2: $w_2(X)$; and T_3: $w_3(X)$:

S_g: $r_1(X)$; $w_2(X)$; $w_1(X)$; $w_3(X)$; c_1; c_2; c_3;

In S_g the operations $w_2(X)$ and $w_3(X)$ are blind writes, since T_2 and T_3 do not read the value of X. The schedule S_g is view serializable, since it is view equivalent to the serial schedule T_1, T_2, T_3. However, S_g is not conflict serializable, since it is not conflict equivalent to any serial schedule. It has been shown that any conflict-serializable schedule is also view serializable but not vice versa, as illustrated by the preceding example. There is an algorithm to test whether a schedule S is view serializable or not. However, the problem of testing for view serializability has been shown to be NP-hard, meaning that finding an efficient polynomial time algorithm for this problem is highly unlikely.

21.5.5 Other Types of Equivalence of Schedules

Serializability of schedules is sometimes considered to be too restrictive as a condition for ensuring the correctness of concurrent executions. Some applications can produce schedules that are correct by satisfying conditions less stringent than either conflict serializability or view serializability. An example is the type of transactions known as **debit-credit transactions**—for example, those that apply deposits and withdrawals to a data item whose value is the current balance of a bank account. The semantics of debit-credit operations is that they update the value of a data item X by either subtracting from or adding to the value of the data item. Because addition and subtraction operations are commutative—that is, they can be applied in any order—it is possible to produce correct schedules that are not serializable. For example, consider the following transactions, each of which may be used to transfer an amount of money between two bank accounts:

T_1: $r_1(X)$; $X := X - 10$; $w_1(X)$; $r_1(Y)$; $Y := Y + 10$; $w_1(Y)$;
T_2: $r_2(Y)$; $Y := Y - 20$; $w_2(Y)$; $r_2(X)$; $X := X + 20$; $w_2(X)$;

Consider the following nonserializable schedule S_h for the two transactions:

S_h: $r_1(X)$; $w_1(X)$; $r_2(Y)$; $w_2(Y)$; $r_1(Y)$; $w_1(Y)$; $r_2(X)$; $w_2(X)$;

With the additional knowledge, or **semantics**, that the operations between each $r_i(I)$ and $w_i(I)$ are commutative, we know that the order of executing the sequences consisting of (read, update, write) is not important as long as each (read, update, write) sequence by a particular transaction T_i on a particular item I is not interrupted by conflicting operations. Hence, the schedule S_h is considered to be correct even though it is not serializable. Researchers have been working on extending concurrency control theory to deal with cases where serializability is considered to be too restrictive as a condition for correctness of schedules. Also, in certain domains of applications such as computer aided design (CAD) of complex systems like aircraft, design transactions last over a long time period. In such applications, more relaxed schemes of concurrency control have been proposed to maintain consistency of the database.

21.6 Transaction Support in SQL

In this section, we give a brief introduction to transaction support in SQL. There are many more details, and the newer standards have more commands for transaction processing. The basic definition of an SQL transaction is similar to our already defined concept of a transaction. That is, it is a logical unit of work and is guaranteed to be atomic. A single SQL statement is always considered to be atomic—either it completes execution without an error or it fails and leaves the database unchanged.

With SQL, there is no explicit Begin_Transaction statement. Transaction initiation is done implicitly when particular SQL statements are encountered. However, every transaction must have an explicit end statement, which is either a COMMIT or a ROLLBACK. Every transaction has certain characteristics attributed to it. These characteristics are specified by a SET TRANSACTION statement in SQL. The characteristics are the *access mode*, the *diagnostic area size*, and the *isolation level*.

The **access mode** can be specified as READ ONLY or READ WRITE. The default is READ WRITE, unless the isolation level of READ UNCOMMITTED is specified (see below), in which case READ ONLY is assumed. A mode of READ WRITE allows select, update, insert, delete, and create commands to be executed. A mode of READ ONLY, as the name implies, is simply for data retrieval.

The **diagnostic area size** option, DIAGNOSTIC SIZE *n*, specifies an integer value *n*, which indicates the number of conditions that can be held simultaneously in the diagnostic area. These conditions supply feedback information (errors or exceptions) to the user or program on the *n* most recently executed SQL statement.

The **isolation level** option is specified using the statement ISOLATION LEVEL <isolation>, where the value for <isolation> can be READ UNCOMMITTED, READ COMMITTED, REPEATABLE READ, or SERIALIZABLE.[15] The default isolation level is SERIALIZABLE, although some systems use READ COMMITTED as their default. The use of the term SERIALIZABLE here is based on not allowing violations that cause dirty read, unrepeatable read, and phantoms,[16] and it is thus not identical to the way serializability was defined earlier in Section 21.5. If a transaction executes at a lower isolation level than SERIALIZABLE, then one or more of the following three violations may occur:

1. **Dirty read.** A transaction T_1 may read the update of a transaction T_2, which has not yet committed. If T_2 fails and is aborted, then T_1 would have read a value that does not exist and is incorrect.

[15]These are similar to the *isolation levels* discussed briefly at the end of Section 21.3.

[16]The dirty read and unrepeatable read problems were discussed in Section 21.1.3. Phantoms are discussed in Section 22.7.1.

transaction performance by relaxing serializability if that is acceptable for their applications.

21.7 Summary

In this chapter we discussed DBMS concepts for transaction processing. We introduced the concept of a database transaction and the operations relevant to transaction processing. We compared single-user systems to multiuser systems and then presented examples of how uncontrolled execution of concurrent transactions in a multiuser system can lead to incorrect results and database values. We also discussed the various types of failures that may occur during transaction execution.

Next we introduced the typical states that a transaction passes through during execution, and discussed several concepts that are used in recovery and concurrency control methods. The system log keeps track of database accesses, and the system uses this information to recover from failures. A transaction either succeeds and reaches its commit point or it fails and has to be rolled back. A committed transaction has its changes permanently recorded in the database. We presented an overview of the desirable properties of transactions—atomicity, consistency preservation, isolation, and durability—which are often referred to as the ACID properties.

Then we defined a schedule (or history) as an execution sequence of the operations of several transactions with possible interleaving. We characterized schedules in terms of their recoverability. Recoverable schedules ensure that, once a transaction commits, it never needs to be undone. Cascadeless schedules add an additional condition to ensure that no aborted transaction requires the cascading abort of other transactions. Strict schedules provide an even stronger condition that allows a simple recovery scheme consisting of restoring the old values of items that have been changed by an aborted transaction.

We defined equivalence of schedules and saw that a serializable schedule is equivalent to some serial schedule. We defined the concepts of conflict equivalence and view equivalence, which led to definitions for conflict serializability and view serializability. A serializable schedule is considered correct. We presented an algorithm for testing the (conflict) serializability of a schedule. We discussed why testing for serializability is impractical in a real system, although it can be used to define and verify concurrency control protocols, and we briefly mentioned less restrictive definitions of schedule equivalence. Finally, we gave a brief overview of how transaction concepts are used in practice within SQL.

Review Questions

21.1. What is meant by the concurrent execution of database transactions in a multiuser system? Discuss why concurrency control is needed, and give informal examples.

21.2. Discuss the different types of failures. What is meant by catastrophic failure?

21.3. Discuss the actions taken by the read_item and write_item operations on a database.

21.4. Draw a state diagram and discuss the typical states that a transaction goes through during execution.

21.5. What is the system log used for? What are the typical kinds of records in a system log? What are transaction commit points, and why are they important?

21.6. Discuss the atomicity, durability, isolation, and consistency preservation properties of a database transaction.

21.7. What is a schedule (history)? Define the concepts of recoverable, cascadeless, and strict schedules, and compare them in terms of their recoverability.

21.8. Discuss the different measures of transaction equivalence. What is the difference between conflict equivalence and view equivalence?

21.9. What is a serial schedule? What is a serializable schedule? Why is a serial schedule considered correct? Why is a serializable schedule considered correct?

21.10. What is the difference between the constrained write and the unconstrained write assumptions? Which is more realistic?

21.11. Discuss how serializability is used to enforce concurrency control in a database system. Why is serializability sometimes considered too restrictive as a measure of correctness for schedules?

21.12. Describe the four levels of isolation in SQL.

21.13. Define the violations caused by each of the following: dirty read, nonrepeatable read, and phantoms.

Exercises

21.14. Change transaction T_2 in Figure 21.2(b) to read

```
read_item(X);
X := X + M;
if X > 90 then exit
else write_item(X);
```

Discuss the final result of the different schedules in Figure 21.3(a) and (b), where $M = 2$ and $N = 2$, with respect to the following questions: Does adding the above condition change the final outcome? Does the outcome obey the implied consistency rule (that the capacity of X is 90)?

21.15. Repeat Exercise 21.14, adding a check in T_1 so that Y does not exceed 90.

21.16. Add the operation commit at the end of each of the transactions T_1 and T_2 in Figure 21.2, and then list all possible schedules for the modified transactions. Determine which of the schedules are recoverable, which are cascadeless, and which are strict.

21.17. List all possible schedules for transactions T_1 and T_2 in Figure 21.2, and determine which are conflict serializable (correct) and which are not.

21.18. How many *serial* schedules exist for the three transactions in Figure 21.8(a)? What are they? What is the total number of possible schedules?

21.19. Write a program to create all possible schedules for the three transactions in Figure 21.8(a), and to determine which of those schedules are conflict serializable and which are not. For each conflict-serializable schedule, your program should print the schedule and list all equivalent serial schedules.

21.20. Why is an explicit transaction end statement needed in SQL but not an explicit begin statement?

21.21. Describe situations where each of the different isolation levels would be useful for transaction processing.

21.22. Which of the following schedules is (conflict) serializable? For each serializable schedule, determine the equivalent serial schedules.

a. $r_1(X); r_3(X); w_1(X); r_2(X); w_3(X);$

b. $r_1(X); r_3(X); w_3(X); w_1(X); r_2(X);$

c. $r_3(X); r_2(X); w_3(X); r_1(X); w_1(X);$

d. $r_3(X); r_2(X); r_1(X); w_3(X); w_1(X);$

21.23. Consider the three transactions T_1, T_2, and T_3, and the schedules S_1 and S_2 given below. Draw the serializability (precedence) graphs for S_1 and S_2, and state whether each schedule is serializable or not. If a schedule is serializable, write down the equivalent serial schedule(s).

T_1: $r_1 (X); r_1 (Z); w_1 (X);$
T_2: $r_2 (Z); r_2 (Y); w_2 (Z); w_2 (Y);$
T_3: $r_3 (X); r_3 (Y); w_3 (Y);$
S_1: $r_1 (X); r_2 (Z); r_1 (Z); r_3 (X); r_3 (Y); w_1 (X); w_3 (Y); r_2 (Y); w_2 (Z); w_2 (Y);$
S_2: $r_1 (X); r_2 (Z); r_3 (X); r_1 (Z); r_2 (Y); r_3 (Y); w_1 (X); w_2 (Z); w_3 (Y); w_2 (Y);$

21.24. Consider schedules S_3, S_4, and S_5 below. Determine whether each schedule is strict, cascadeless, recoverable, or nonrecoverable. (Determine the strictest recoverability condition that each schedule satisfies.)

S_3: $r_1 (X); r_2 (Z); r_1 (Z); r_3 (X); r_3 (Y); w_1 (X); c_1; w_3 (Y); c_3; r_2 (Y); w_2 (Z); w_2 (Y); c_2;$
S_4: $r_1 (X); r_2 (Z); r_1 (Z); r_3 (X); r_3 (Y); w_1 (X); w_3 (Y); r_2 (Y); w_2 (Z); w_2 (Y); c_1; c_2; c_3;$
S_5: $r_1 (X); r_2 (Z); r_3 (X); r_1 (Z); r_2 (Y); r_3 (Y); w_1 (X); c_1; w_2 (Z); w_3 (Y); w_2 (Y); c_3; c_2;$

chapter 22

Concurrency Control Techniques

In this chapter we discuss a number of concurrency control techniques that are used to ensure the noninterference or isolation property of concurrently executing transactions. Most of these techniques ensure serializability of schedules—which we defined in Section 21.5—using **concurrency control protocols** (sets of rules) that guarantee serializability. One important set of protocols—known as *two-phase locking protocols*—employ the technique of **locking** data items to prevent multiple transactions from accessing the items concurrently; a number of locking protocols are described in Sections 22.1 and 22.3.2. Locking protocols are used in most commercial DBMSs. Another set of concurrency control protocols use **timestamps**. A timestamp is a unique identifier for each transaction, generated by the system. Timestamps values are generated in the same order as the transaction start times. Concurrency control protocols that use timestamp ordering to ensure serializability are introduced in Section 22.2. In Section 22.3 we discuss **multiversion** concurrency control protocols that use multiple versions of a data item. One multiversion protocol extends timestamp order to multiversion timestamp ordering (Section 22.3.1), and another extends two-phase locking (Section 22.3.2). In Section 22.4 we present a protocol based on the concept of **validation** or **certification** of a transaction after it executes its operations; these are sometimes called **optimistic protocols**, and also assume that multiple versions of a data item can exist.

Another factor that affects concurrency control is the **granularity** of the data items—that is, what portion of the database a data item represents. An item can be as small as a single attribute (field) value or as large as a disk block, or even a whole file or the entire database. We discuss granularity of items and a multiple granularity concurrency control protocol, which is an extension of two-phase locking, in Section 22.5. In Section 22.6 we describe concurrency control issues that arise when indexes are used to process transactions, and in Section 22.7 we discuss some additional concurrency control concepts. Section 22.8 summarizes the chapter.

It is sufficient to read Sections 22.1, 22.5, 22.6, and 22.7, and possibly 22.3.2, if your main interest is an introduction to the concurrency control techniques that are based on locking, which are used most often in practice. The other techniques are mainly of theoretical interest.

22.1 Two-Phase Locking Techniques for Concurrency Control

Some of the main techniques used to control concurrent execution of transactions are based on the concept of locking data items. A **lock** is a variable associated with a data item that describes the status of the item with respect to possible operations that can be applied to it. Generally, there is one lock for each data item in the database. Locks are used as a means of synchronizing the access by concurrent transactions to the database items. In Section 22.1.1 we discuss the nature and types of locks. Then, in Section 22.1.2 we present protocols that use locking to guarantee serializability of transaction schedules. Finally, in Section 22.1.3 we describe two problems associated with the use of locks—deadlock and starvation—and show how these problems are handled in concurrency control protocols.

22.1.1 Types of Locks and System Lock Tables

Several types of locks are used in concurrency control. To introduce locking concepts gradually, first we discuss binary locks, which are simple, but are also *too restrictive for database concurrency control purposes*, and so are not used in practice. Then we discuss *shared/exclusive* locks—also known as *read/write* locks—which provide more general locking capabilities and are used in practical database locking schemes. In Section 22.3.2 we describe an additional type of lock called a *certify lock*, and show how it can be used to improve performance of locking protocols.

Binary Locks. A **binary lock** can have two **states** or **values:** locked and unlocked (or 1 and 0, for simplicity). A distinct lock is associated with each database item X. If the value of the lock on X is 1, item X *cannot be accessed* by a database operation that requests the item. If the value of the lock on X is 0, the item can be accessed when requested, and the lock value is changed to 1. We refer to the current value (or state) of the lock associated with item X as **lock**($\boldsymbol{X}$).

Two operations, lock_item and unlock_item, are used with binary locking. A transaction requests access to an item X by first issuing a **lock_item**($\boldsymbol{X}$) operation. If LOCK(X) = 1, the transaction is forced to wait. If LOCK(X) = 0, it is set to 1 (the transaction **locks** the item) and the transaction is allowed to access item X. When the transaction is through using the item, it issues an **unlock_item**($\boldsymbol{X}$) operation, which sets LOCK(X) back to 0 (**unlocks** the item) so that X may be accessed by other transactions. Hence, a binary lock enforces **mutual exclusion** on the data item. A description of the lock_item(X) and unlock_item(X) operations is shown in Figure 22.1.

```
lock_item(X):
B:  if LOCK(X) = 0                  (* item is unlocked *)
        then LOCK(X) ←1             (* lock the item *)
    else
        begin
        wait (until LOCK(X) = 0
            and the lock manager wakes up the transaction);
        go to B
        end;
unlock_item(X):
    LOCK(X) ← 0;                        (* unlock the item *)
    if any transactions are waiting
        then wakeup one of the waiting transactions;
```

Figure 22.1
Lock and unlock operations for binary locks.

Notice that the lock_item and unlock_item operations must be implemented as indivisible units (known as **critical sections** in operating systems); that is, no interleaving should be allowed once a lock or unlock operation is started until the operation terminates or the transaction waits. In Figure 22.1, the wait command within the lock_item(X) operation is usually implemented by putting the transaction in a waiting queue for item X until X is unlocked and the transaction can be granted access to it. Other transactions that also want to access X are placed in the same queue. Hence, the wait command is considered to be outside the lock_item operation.

It is quite simple to implement a binary lock; all that is needed is a binary-valued variable, LOCK, associated with each data item X in the database. In its simplest form, each lock can be a record with three fields: <Data_item_name, LOCK, Locking_transaction> plus a queue for transactions that are waiting to access the item. The system needs to maintain *only these records for the items that are currently locked* in a **lock table**, which could be organized as a hash file on the item name. Items not in the lock table are considered to be unlocked. The DBMS has a **lock manager subsystem** to keep track of and control access to locks.

If the simple binary locking scheme described here is used, every transaction must obey the following rules:

1. A transaction T must issue the operation lock_item(X) before any read_item(X) or write_item(X) operations are performed in T.
2. A transaction T must issue the operation unlock_item(X) after all read_item(X) and write_item(X) operations are completed in T.
3. A transaction T will not issue a lock_item(X) operation if it already holds the lock on item X.[1]
4. A transaction T will not issue an unlock_item(X) operation unless it already holds the lock on item X.

[1]This rule may be removed if we modify the lock_item (X) operation in Figure 22.1 so that if the item is currently locked *by the requesting transaction*, the lock is granted.

These rules can be enforced by the lock manager module of the DBMS. Between the lock_item(X) and unlock_item(X) operations in transaction T, T is said to **hold the lock** on item X. At most one transaction can hold the lock on a particular item. Thus no two transactions can access the same item concurrently.

Shared/Exclusive (or Read/Write) Locks. The preceding binary locking scheme is too restrictive for database items because at most, one transaction can hold a lock on a given item. We should allow several transactions to access the same item X if they all access X for *reading purposes only*. This is because read operations on the same item by different transactions are not conflicting (see Section 21.4.1). However, if a transaction is to write an item X, it must have exclusive access to X. For this purpose, a different type of lock called a **multiple-mode lock** is used. In this scheme—called **shared/exclusive** or **read/write** locks—there are three locking operations: read_lock(X), write_lock(X), and unlock(X). A lock associated with an item X, LOCK(X), now has three possible states: *read-locked*, *write-locked*, or *unlocked*. A **read-locked item** is also called **share-locked** because other transactions are allowed to read the item, whereas a **write-locked item** is called **exclusive-locked** because a single transaction exclusively holds the lock on the item.

One method for implementing the preceding operations on a read/write lock is to keep track of the number of transactions that hold a shared (read) lock on an item in the lock table. Each record in the lock table will have four fields: <Data_item_name, LOCK, No_of_reads, Locking_transaction(s)>. Again, to save space, the system needs to maintain lock records only for locked items in the lock table. The value (state) of LOCK is either read-locked or write-locked, suitably coded (if we assume no records are kept in the lock table for unlocked items). If LOCK(X)=write-locked, the value of locking_transaction(s) is a single transaction that holds the exclusive (write) lock on X. If LOCK(X)=read-locked, the value of locking transaction(s) is a list of one or more transactions that hold the shared (read) lock on X. The three operations read_lock(X), write_lock(X), and unlock(X) are described in Figure 22.2.[2] As before, each of the three locking operations should be considered indivisible; no interleaving should be allowed once one of the operations is started until either the operation terminates by granting the lock or the transaction is placed in a waiting queue for the item.

When we use the shared/exclusive locking scheme, the system must enforce the following rules:

1. A transaction T must issue the operation read_lock(X) or write_lock(X) before any read_item(X) operation is performed in T.
2. A transaction T must issue the operation write_lock(X) before any write_item(X) operation is performed in T.

---

[2]These algorithms do not allow *upgrading* or *downgrading* of locks, as described later in this section. The reader can extend the algorithms to allow these additional operations.

```
read_lock(X):
B:  if LOCK(X) = "unlocked"
        then begin LOCK(X) ← "read-locked";
                no_of_reads(X) ← 1
                end
    else if LOCK(X) = "read-locked"
        then no_of_reads(X) ← no_of_reads(X) + 1
    else begin
            wait (until LOCK(X) = "unlocked"
                and the lock manager wakes up the transaction);
            go to B
            end;
write_lock(X):
B:  if LOCK(X) = "unlocked"
        then LOCK(X) ← "write-locked"
    else begin
            wait (until LOCK(X) = "unlocked"
                and the lock manager wakes up the transaction);
            go to B
            end;
unlock (X):
    if LOCK(X) = "write-locked"
        then begin LOCK(X) ← "unlocked";
                wakeup one of the waiting transactions, if any
                end
    else it LOCK(X) = "read-locked"
        then begin
                no_of_reads(X) ← no_of_reads(X) −1;
                if no_of_reads(X) = 0
                    then begin LOCK(X) = "unlocked";
                            wakeup one of the waiting transactions, if any
                            end
                end;
```

Figure 22.2
Locking and unlocking operations for two-mode (read-write or shared-exclusive) locks.

3. A transaction T must issue the operation unlock(X) after all read_item(X) and write_item(X) operations are completed in T.[3]
4. A transaction T will not issue a read_lock(X) operation if it already holds a read (shared) lock or a write (exclusive) lock on item X. This rule may be relaxed, as we discuss shortly.

[3]This rule may be relaxed to allow a transaction to unlock an item, then lock it again later.

5. A transaction T will not issue a write_lock(X) operation if it already holds a read (shared) lock or write (exclusive) lock on item X. This rule may also be relaxed, as we discuss shortly.
6. A transaction T will not issue an unlock(X) operation unless it already holds a read (shared) lock or a write (exclusive) lock on item X.

Conversion of Locks. Sometimes it is desirable to relax conditions 4 and 5 in the preceding list in order to allow **lock conversion**; that is, a transaction that already holds a lock on item X is allowed under certain conditions to **convert** the lock from one locked state to another. For example, it is possible for a transaction T to issue a read_lock(X) and then later to **upgrade** the lock by issuing a write_lock(X) operation. If T is the only transaction holding a read lock on X at the time it issues the write_lock(X) operation, the lock can be upgraded; otherwise, the transaction must wait. It is also possible for a transaction T to issue a write_lock(X) and then later to **downgrade** the lock by issuing a read_lock(X) operation. When upgrading and downgrading of locks is used, the lock table must include transaction identifiers in the record structure for each lock (in the locking_transaction(s) field) to store the information on which transactions hold locks on the item. The descriptions of the read_lock(X) and write_lock(X) operations in Figure 22.2 must be changed appropriately to allow for lock upgrading and downgrading. We leave this as an exercise for the reader.

Using binary locks or read/write locks in transactions, as described earlier, does not guarantee serializability of schedules on its own. Figure 22.3 shows an example where the preceding locking rules are followed but a nonserializable schedule may result. This is because in Figure 22.3(a) the items Y in T_1 and X in T_2 were unlocked too early. This allows a schedule such as the one shown in Figure 22.3(c) to occur, which is not a serializable schedule and hence gives incorrect results. To guarantee serializability, we must follow *an additional protocol* concerning the positioning of locking and unlocking operations in every transaction. The best-known protocol, two-phase locking, is described in the next section.

22.1.2 Guaranteeing Serializability by Two-Phase Locking

A transaction is said to follow the **two-phase locking protocol** if *all* locking operations (read_lock, write_lock) precede the *first* unlock operation in the transaction.[4] Such a transaction can be divided into two phases: an **expanding** or **growing (first) phase**, during which new locks on items can be acquired but none can be released; and a **shrinking (second) phase**, during which existing locks can be released but no new locks can be acquired. If lock conversion is allowed, then upgrading of locks (from read-locked to write-locked) must be done during the expanding phase, and downgrading of locks (from write-locked to read-locked) must be done in the

[4]This is unrelated to the two-phase commit protocol for recovery in distributed databases (see Chapter 25).

(a)

T_1	T_2
read_lock(Y);	read_lock(X);
read_item(Y);	read_item(X);
unlock(Y);	unlock(X);
write_lock(X);	write_lock(Y);
read_item(X);	read_item(Y);
$X := X + Y$;	$Y := X + Y$;
write_item(X);	write_item(Y);
unlock(X);	unlock(Y);

(b) Initial values: X=20, Y=30

Result serial schedule T_1 followed by T_2: X=50, Y=80

Result of serial schedule T_2 followed by T_1: X=70, Y=50

(c)

Time

T_1	T_2
read_lock(Y);	
read_item(Y);	
unlock(Y);	
	read_lock(X);
	read_item(X);
	unlock(X);
	write_lock(Y);
	read_item(Y);
	$Y := X + Y$;
	write_item(Y);
	unlock(Y);
write_lock(X);	
read_item(X);	
$X := X + Y$;	
write_item(X);	
unlock(X);	

Result of schedule S:
X=50, Y=50
(nonserializable)

Figure 22.3
Transactions that do not obey two-phase locking. (a) Two transactions T_1 and T_2. (b) Results of possible serial schedules of T_1 and T_2. (c) A nonserializable schedule S that uses locks.

shrinking phase. Hence, a read_lock(X) operation that downgrades an already held write lock on X can appear only in the shrinking phase.

Transactions T_1 and T_2 in Figure 22.3(a) do not follow the two-phase locking protocol because the write_lock(X) operation follows the unlock(Y) operation in T_1, and similarly the write_lock(Y) operation follows the unlock(X) operation in T_2. If we enforce two-phase locking, the transactions can be rewritten as T_1' and T_2', as shown in Figure 22.4. Now, the schedule shown in Figure 22.3(c) is not permitted for T_1' and T_2' (with their modified order of locking and unlocking operations) under the rules of locking described in Section 22.1.1 because T_1' will issue its write_lock(X) *before* it unlocks item Y; consequently, when T_2' issues its read_lock(X), it is forced to wait until T_1' releases the lock by issuing an unlock (X) in the schedule.

T_1'	T_2'

read_lock(Y);
read_item(Y);
write_lock(X);
unlock(Y)
read_item(X);
$X := X + Y$;
write_item(X);
unlock(X);
	read_lock(X);
	read_item(X);
	write_lock(Y);
	unlock(X)
	read_item(Y);
	$Y := X + Y$;
	write_item(Y);
	unlock(Y);

Figure 22.4
Transactions T_1' and T_2', which are the same as T_1 and T_2 in Figure 22.3, but follow the two-phase locking protocol. Note that they can produce a deadlock.

It can be proved that, if *every* transaction in a schedule follows the two-phase locking protocol, the schedule is *guaranteed to be serializable*, obviating the need to test for serializability of schedules. The locking protocol, by enforcing two-phase locking rules, also enforces serializability.

Two-phase locking may limit the amount of concurrency that can occur in a schedule because a transaction T may not be able to release an item X after it is through using it if T must lock an additional item Y later; or conversely, T must lock the additional item Y before it needs it so that it can release X. Hence, X must remain locked by T until all items that the transaction needs to read or write have been locked; only then can X be released by T. Meanwhile, another transaction seeking to access X may be forced to wait, even though T is done with X; conversely, if Y is locked earlier than it is needed, another transaction seeking to access Y is forced to wait even though T is not using Y yet. This is the price for guaranteeing serializability of all schedules without having to check the schedules themselves.

Although the two-phase locking protocol guarantees serializability (that is, every schedule that is permitted is serializable), it does not permit *all possible* serializable schedules (that is, some serializable schedules will be prohibited by the protocol).

Basic, Conservative, Strict, and Rigorous Two-Phase Locking. There are a number of variations of two-phase locking (2PL). The technique just described is known as **basic 2PL**. A variation known as **conservative 2PL** (or **static 2PL**) requires a transaction to lock all the items it accesses *before the transaction begins execution*, by **predeclaring** its *read-set* and *write-set*. Recall from Section 21.1.2 that the **read-set** of a transaction is the set of all items that the transaction reads, and the **write-set** is the set of all items that it writes. If any of the predeclared items needed cannot be locked, the transaction does not lock any item; instead, it waits until all the items are available for locking. Conservative 2PL is a deadlock-free protocol, as we will see in Section 22.1.3 when we discuss the deadlock problem. However, it is difficult to use in practice because of the need to predeclare the read-set and write-set, which is not possible in many situations.

In practice, the most popular variation of 2PL is **strict 2PL**, which guarantees strict schedules (see Section 21.4). In this variation, a transaction T does not release any of its exclusive (write) locks until *after* it commits or aborts. Hence, no other transaction can read or write an item that is written by T unless T has committed, leading to a strict schedule for recoverability. Strict 2PL is not deadlock-free. A more restrictive variation of strict 2PL is **rigorous 2PL**, which also guarantees strict schedules. In this variation, a transaction T does not release any of its locks (exclusive or shared) until after it commits or aborts, and so it is easier to implement than strict 2PL. Notice the difference between conservative and rigorous 2PL: the former must lock all its items *before it starts*, so once the transaction starts it is in its shrinking phase; the latter does not unlock any of its items until *after it terminates* (by committing or aborting), so the transaction is in its expanding phase until it ends.

In many cases, the **concurrency control subsystem** itself is responsible for generating the read_lock and write_lock requests. For example, suppose the system is to enforce the strict 2PL protocol. Then, whenever transaction T issues a read_item(X), the system calls the read_lock(X) operation on behalf of T. If the state of LOCK(X) is write_locked by some other transaction T', the system places T in the waiting queue for item X; otherwise, it grants the read_lock(X) request and permits the read_item(X) operation of T to execute. On the other hand, if transaction T issues a write_item(X), the system calls the write_lock(X) operation on behalf of T. If the state of LOCK(X) is write_locked or read_locked by some other transaction T', the system places T in the waiting queue for item X; if the state of LOCK(X) is read_locked and T itself is the only transaction holding the read lock on X, the system upgrades the lock to write_locked and permits the write_item(X) operation by T. Finally, if the state of LOCK(X) is unlocked, the system grants the write_lock(X) request and permits the write_item(X) operation to execute. After each action, the system must update its lock table appropriately.

The use of locks can cause two additional problems: deadlock and starvation. We discuss these problems and their solutions in the next section.

22.1.3 Dealing with Deadlock and Starvation

Deadlock occurs when *each* transaction T in a set of *two or more transactions* is waiting for some item that is locked by some other transaction T' in the set. Hence, each transaction in the set is in a waiting queue, waiting for one of the other transactions in the set to release the lock on an item. But because the other transaction is also waiting, it will never release the lock. A simple example is shown in Figure 22.5(a), where the two transactions T_1' and T_2' are deadlocked in a partial schedule; T_1' is in the waiting queue for X, which is locked by T_2', while T_2' is in the waiting queue for Y, which is locked by T_1'. Meanwhile, neither T_1' nor T_2' nor any other transaction can access items X and Y.

Deadlock Prevention Protocols. One way to prevent deadlock is to use a **deadlock prevention protocol**.[5] One deadlock prevention protocol, which is used

[5]These protocols are not generally used in practice, either because of unrealistic assumptions or because of their possible overhead. Deadlock detection and timeouts (covered in the following sections) are more practical.

(a)

Time ↓

T_1'	T_2'
read_lock(Y);	
read_item(Y);	
	read_lock(X);
	read_item(X);
write_lock(X);	
	write_lock(Y);

(b)

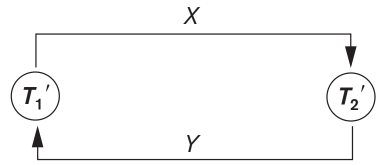


Figure 22.5
Illustrating the deadlock problem. (a) A partial schedule of T_1' and T_2' that is in a state of deadlock. (b) A wait-for graph for the partial schedule in (a).

in conservative two-phase locking, requires that every transaction lock *all the items it needs in advance* (which is generally not a practical assumption)—if any of the items cannot be obtained, none of the items are locked. Rather, the transaction waits and then tries again to lock all the items it needs. Obviously this solution further limits concurrency. A second protocol, which also limits concurrency, involves *ordering all the items* in the database and making sure that a transaction that needs several items will lock them according to that order. This requires that the programmer (or the system) is aware of the chosen order of the items, which is also not practical in the database context.

A number of other deadlock prevention schemes have been proposed that make a decision about what to do with a transaction involved in a possible deadlock situation: Should it be blocked and made to wait or should it be aborted, or should the transaction preempt and abort another transaction? Some of these techniques use the concept of **transaction timestamp** TS(T), which is a unique identifier assigned to each transaction. The timestamps are typically based on the order in which transactions are started; hence, if transaction T_1 starts before transaction T_2, then TS(T_1) < TS(T_2). Notice that the *older* transaction (which starts first) has the *smaller* timestamp value. Two schemes that prevent deadlock are called *wait-die* and *wound-wait*. Suppose that transaction T_i tries to lock an item X but is not able to because X is locked by some other transaction T_j with a conflicting lock. The rules followed by these schemes are:

- **Wait-die.** If TS(T_i) < TS(T_j), then (T_i older than T_j) T_i is allowed to wait; otherwise (T_i younger than T_j) abort T_i (T_i *dies*) and restart it later *with the same timestamp.*
- **Wound-wait.** If TS(T_i) < TS(T_j), then (T_i older than T_j) abort T_j (T_i *wounds* T_j) and restart it later *with the same timestamp;* otherwise (T_i younger than T_j) T_i is allowed to wait.

In wait-die, an older transaction is allowed to *wait for a younger transaction*, whereas a younger transaction requesting an item held by an older transaction is aborted and restarted. The wound-wait approach does the opposite: A younger transaction is allowed to *wait for an older one*, whereas an older transaction requesting an item held by a younger transaction *preempts* the younger transaction by aborting it. Both schemes end up aborting the *younger* of the two transactions (the transaction that started later) that *may be involved* in a deadlock, assuming that this will waste less processing. It can be shown that these two techniques are *deadlock-free*, since in wait-die, transactions only wait for younger transactions so no cycle is created. Similarly, in wound-wait, transactions only wait for older transactions so no cycle is created. However, both techniques may cause some transactions to be aborted and restarted needlessly, even though those transactions may *never actually cause a deadlock*.

Another group of protocols that prevent deadlock do not require timestamps. These include the no waiting (NW) and cautious waiting (CW) algorithms. In the **no waiting algorithm**, if a transaction is unable to obtain a lock, it is immediately aborted and then restarted after a certain time delay without checking whether a deadlock will actually occur or not. In this case, no transaction ever waits, so no deadlock will occur. However, this scheme can cause transactions to abort and restart needlessly. The **cautious waiting** algorithm was proposed to try to reduce the number of needless aborts/restarts. Suppose that transaction T_i tries to lock an item X but is not able to do so because X is locked by some other transaction T_j with a conflicting lock. The cautious waiting rules are as follows:

- **Cautious waiting.** If T_j is not blocked (not waiting for some other locked item), then T_i is blocked and allowed to wait; otherwise abort T_i.

It can be shown that cautious waiting is deadlock-free, because no transaction will ever wait for another blocked transaction. By considering the time $b(T)$ at which each blocked transaction T was blocked, if the two transactions T_i and T_j above both become blocked, and T_i is waiting for T_j, then $b(T_i) < b(T_j)$, since T_i can only wait for T_j at a time when T_j is not blocked itself. Hence, the blocking times form a total ordering on all blocked transactions, so no cycle that causes deadlock can occur.

Deadlock Detection. A second, more practical approach to dealing with deadlock is **deadlock detection**, where the system checks if a state of deadlock actually exists. This solution is attractive if we know there will be little interference among the transactions—that is, if different transactions will rarely access the same items at the same time. This can happen if the transactions are short and each transaction locks only a few items, or if the transaction load is light. On the other hand, if transactions are long and each transaction uses many items, or if the transaction load is quite heavy, it may be advantageous to use a deadlock prevention scheme.

A simple way to detect a state of deadlock is for the system to construct and maintain a **wait-for graph**. One node is created in the wait-for graph for each transaction that is currently executing. Whenever a transaction T_i is waiting to lock an item X that is currently locked by a transaction T_j, a directed edge ($T_i \rightarrow T_j$) is created in the wait-for graph. When T_j releases the lock(s) on the items that T_i was waiting for, the directed edge is dropped from the wait-for graph. We have a state of deadlock if and only if the wait-for graph has a cycle. One problem with this approach is the matter of determining *when* the system should check for a deadlock. One possibility is to check for a cycle every time an edge is added to the wait-for graph, but this may cause excessive overhead. Criteria such as the number of currently executing transactions or the period of time several transactions have been waiting to lock items may be used instead to check for a cycle. Figure 22.5(b) shows the wait-for graph for the (partial) schedule shown in Figure 22.5(a).

If the system is in a state of deadlock, some of the transactions causing the deadlock must be aborted. Choosing which transactions to abort is known as **victim selection**. The algorithm for victim selection should generally avoid selecting transactions that have been running for a long time and that have performed many updates, and it should try instead to select transactions that have not made many changes (younger transactions).

Timeouts. Another simple scheme to deal with deadlock is the use of **timeouts**. This method is practical because of its low overhead and simplicity. In this method, if a transaction waits for a period longer than a system-defined timeout period, the system assumes that the transaction may be deadlocked and aborts it—regardless of whether a deadlock actually exists or not.

Starvation. Another problem that may occur when we use locking is **starvation**, which occurs when a transaction cannot proceed for an indefinite period of time while other transactions in the system continue normally. This may occur if the waiting scheme for locked items is unfair, giving priority to some transactions over others. One solution for starvation is to have a fair waiting scheme, such as using a **first-come-first-served** queue; transactions are enabled to lock an item in the order in which they originally requested the lock. Another scheme allows some transactions to have priority over others but increases the priority of a transaction the longer it waits, until it eventually gets the highest priority and proceeds. Starvation can also occur because of victim selection if the algorithm selects the same transaction as victim repeatedly, thus causing it to abort and never finish execution. The algorithm can use higher priorities for transactions that have been aborted multiple times to avoid this problem. The wait-die and wound-wait schemes discussed previously avoid starvation, because they restart a transaction that has been aborted with its same original timestamp, so the possibility that the same transaction is aborted repeatedly is slim.

22.2 Concurrency Control Based on Timestamp Ordering

The use of locks, combined with the 2PL protocol, guarantees serializability of schedules. The serializable schedules produced by 2PL have their equivalent serial schedules based on the order in which executing transactions lock the items they acquire. If a transaction needs an item that is already locked, it may be forced to wait until the item is released. Some transactions may be aborted and restarted because of the deadlock problem. A different approach that guarantees serializability involves using transaction timestamps to order transaction execution for an equiva-

Review Questions

22.1. What is the two-phase locking protocol? How does it guarantee serializability?

22.2. What are some variations of the two-phase locking protocol? Why is strict or rigorous two-phase locking often preferred?

22.3. Discuss the problems of deadlock and starvation, and the different approaches to dealing with these problems.

22.4. Compare binary locks to exclusive/shared locks. Why is the latter type of locks preferable?

22.5. Describe the wait-die and wound-wait protocols for deadlock prevention.

22.6. Describe the cautious waiting, no waiting, and timeout protocols for deadlock prevention.

22.7. What is a timestamp? How does the system generate timestamps?

22.8. Discuss the timestamp ordering protocol for concurrency control. How does strict timestamp ordering differ from basic timestamp ordering?

22.9. Discuss two multiversion techniques for concurrency control.

22.10. What is a certify lock? What are the advantages and disadvantages of using certify locks?

22.11. How do optimistic concurrency control techniques differ from other concurrency control techniques? Why are they also called validation or certification techniques? Discuss the typical phases of an optimistic concurrency control method.

22.12. How does the granularity of data items affect the performance of concurrency control? What factors affect selection of granularity size for data items?

22.13. What type of lock is needed for insert and delete operations?

22.14. What is multiple granularity locking? Under what circumstances is it used?

22.15. What are intention locks?

22.16. When are latches used?

22.17. What is a phantom record? Discuss the problem that a phantom record can cause for concurrency control.

22.18. How does index locking resolve the phantom problem?

22.19. What is a predicate lock?

Exercises

22.20. Prove that the basic two-phase locking protocol guarantees conflict serializability of schedules. (*Hint*: Show that if a serializability graph for a schedule has a cycle, then at least one of the transactions participating in the schedule does not obey the two-phase locking protocol.)

22.21. Modify the data structures for multiple-mode locks and the algorithms for read_lock(X), write_lock(X), and unlock(X) so that upgrading and downgrading of locks are possible. (*Hint*: The lock needs to check the transaction id(s) that hold the lock, if any.)

22.22. Prove that strict two-phase locking guarantees strict schedules.

22.23. Prove that the wait-die and wound-wait protocols avoid deadlock and starvation.

22.24. Prove that cautious waiting avoids deadlock.

22.25. Apply the timestamp ordering algorithm to the schedules in Figure 21.8(b) and (c), and determine whether the algorithm will allow the execution of the schedules.

22.26. Repeat Exercise 22.25, but use the multiversion timestamp ordering method.

22.27. Why is two-phase locking not used as a concurrency control method for indexes such as B^+-trees?

22.28. The compatibility matrix in Figure 22.8 shows that IS and IX locks are compatible. Explain why this is valid.

22.29. The MGL protocol states that a transaction T can unlock a node N, only if none of the children of node N are still locked by transaction T. Show that without this condition, the MGL protocol would be incorrect.

Selected Bibliography

The two-phase locking protocol and the concept of predicate locks were first proposed by Eswaran et al. (1976). Bernstein et al. (1987), Gray and Reuter (1993), and Papadimitriou (1986) focus on concurrency control and recovery. Kumar (1996) focuses on performance of concurrency control methods. Locking is discussed in Gray et al. (1975), Lien and Weinberger (1978), Kedem and Silbershatz (1980), and Korth (1983). Deadlocks and wait-for graphs were formalized by Holt (1972), and the wait-wound and wound-die schemes are presented in Rosenkrantz et al. (1978). Cautious waiting is discussed in Hsu and Zhang (1992). Helal et al. (1993) compares various locking approaches. Timestamp-based concurrency control techniques are discussed in Bernstein and Goodman (1980) and Reed (1983). Optimistic concurrency control is discussed in Kung and Robinson (1981) and Bassiouni (1988). Papadimitriou and Kanellakis (1979) and Bernstein and