SRM Institute of Science and Technology

18CSC303J

Database Management System

Unit- V

Topics covered in Unit 5

- Transaction concepts
- Properties of Transactions
- Serializability
- Testing for Serializability
- System Recovery
- Concurrency control
- Two-phase commit protocol
- Recovery and Atomicity
- Log based recovery
- Concurrency problems
- Locking mechanism
- Deadlock
- Two phase locking protocol
- Isolation
- Intent locking

Transactions

Transaction Concept

- A **transaction** is a *unit* of program execution that accesses and possibly updates various data items.
- E.g., transaction to transfer $50 from account A to account B:
  1. **read**(*A*)
  2. *A* := *A* – 50
  3. **write**(*A*)
  4. **read**(*B*)
  5. *B* := *B* + 50
  6. **write**(*B*)
- Two main issues to deal with:
  - Failures of various kinds, such as hardware failures and system crashes
  - Concurrent execution of multiple transactions

Required Properties of a Transaction

- Consider a transaction to transfer $50 from account A to account B:
  1. **read**(*A*)
  2. *A* := *A* – 50
  3. **write**(*A*)
  4. **read**(*B*)
  5. *B* := *B* + 50
  6. **write**(*B*)
- **Atomicity requirement**
  - If the transaction fails after step 3 and before step 6, money will be "lost" leading to an inconsistent database state
    - Failure could be due to software or hardware
  - The system should ensure that updates of a partially executed transaction are not reflected in the database
- **Durability requirement** — once the user has been notified that the transaction has completed (i.e., the transfer of the $50 has taken place), the updates to the database by the transaction must persist even if there are software or hardware failures.

Required Properties of a Transaction (Cont.)

- **Consistency requirement** in above example:
  - The sum of A and B is unchanged by the execution of the transaction
- In general, consistency requirements include
  - Explicitly specified integrity constraints such as primary keys and foreign keys
  - Implicit integrity constraints
    - e.g., sum of balances of all accounts, minus sum of loan amounts must equal value of cash-in-hand
- A transaction, when starting to execute, must see a consistent database.
- During transaction execution the database may be temporarily inconsistent.
- When the transaction completes successfully the database must be consistent
  - Erroneous transaction logic can lead to inconsistency

Required Properties of a Transaction (Cont.)

- **Isolation requirement** — if between steps 3 and 6 (of the fund transfer transaction) , another transaction **T2** is allowed to access the partially updated database, it will see an inconsistent database (the sum $A + B$ will be less than it should be).

T1	**T2**
1. **read**(A)	
2. $A := A - 50$	
3. **write**(A)	
	read(A), read(B), print(A+B)
4. **read**(B)	
5. $B := B + 50$	
6. **write**(B)	

- Isolation can be ensured trivially by running transactions **serially**
  - That is, one after the other.
- However, executing multiple transactions concurrently has significant benefits, as we will see later.

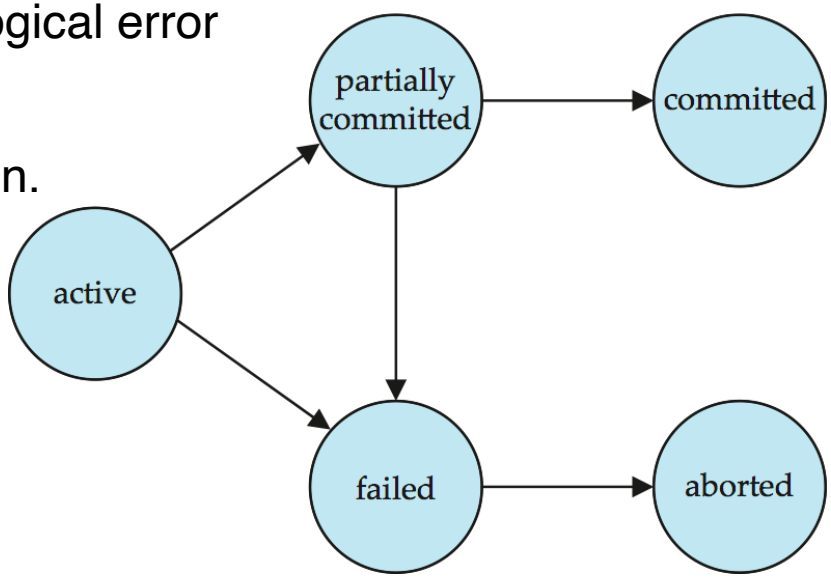
ACID Properties

A **transaction** is a unit of program execution that accesses and possibly updates various data items. To preserve the integrity of data the database system must ensure:

- **Atomicity.** Either all operations of the transaction are properly reflected in the database or none are.
- **Consistency.** Execution of a transaction in isolation preserves the consistency of the database.
- **Isolation.** Although multiple transactions may execute concurrently, each transaction must be unaware of other concurrently executing transactions. Intermediate transaction results must be hidden from other concurrently executed transactions.
  - That is, for every pair of transactions T_i and T_j, it appears to T_i that either T_j, finished execution before T_i started, or T_j started execution after T_i finished.
- **Durability.** After a transaction completes successfully, the changes it has made to the database persist, even if there are system failures.

Transaction State

- **Active** – the initial state; the transaction stays in this state while it is executing
- **Partially committed** – after the final statement has been executed.
- **Failed** -- after the discovery that normal execution can no longer proceed.
- **Aborted** – after the transaction has been rolled back and the database restored to its state prior to the start of the transaction. Two options after it has been aborted:
  - Restart the transaction
    - can be done only if no internal logical error
  - Kill the transaction
- **Committed** – after successful completion.



Concurrent Executions

- Multiple transactions are allowed to run concurrently in the system. Advantages are:
  - **Increased processor and disk utilization**, leading to better transaction *throughput*
    - E.g. one transaction can be using the CPU while another is reading from or writing to the disk
  - **Reduced average response time** for transactions: short transactions need not wait behind long ones.
- **Concurrency control schemes** – mechanisms to achieve isolation
  - That is, to control the interaction among the concurrent transactions in order to prevent them from destroying the consistency of the database

Schedules

- **Schedule** – a sequences of instructions that specify the chronological order in which instructions of concurrent transactions are executed
  - A schedule for a set of transactions must consist of all instructions of those transactions
  - Must preserve the order in which the instructions appear in each individual transaction.
- A transaction that successfully completes its execution will have a **commit** instructions as the last statement
  - By default transaction assumed to execute commit instruction as its last step
- A transaction that fails to successfully complete its execution will have an **abort** instruction as the last statement

Schedule 1

- Let T_1 transfer \$50 from A to B, and T_2 transfer 10% of the balance from A to B.
- An example of a **serial** schedule in which T_1 is followed by T_2 :

T_1	T_2
read (A)	
$A := A - 50$	
write (A)	
read (B)	
$B := B + 50$	
write (B)	
commit	
	read (A)
	$temp := A * 0.1$
	$A := A - temp$
	write (A)
	read (B)
	$B := B + temp$
	write (B)
	commit

Schedule 2

- A **serial** schedule in which T_2 is followed by T_1 :

T_1	T_2
	read (A) $temp := A * 0.1$ $A := A - temp$ write (A) read (B) $B := B + temp$ write (B) commit
read (A) $A := A - 50$ write (A) read (B) $B := B + 50$ write (B) commit	

Schedule 3

- Let T_1 and T_2 be the transactions defined previously. The following schedule is not a serial schedule, but it is **equivalent** to Schedule 1.

T_1	T_2
read (A)	
$A := A - 50$	
write (A)	
	read (A)
	$temp := A * 0.1$
	$A := A - temp$
	write (A)
read (B)	
$B := B + 50$	
write (B)	
commit	
	read (B)
	$B := B + temp$
	write (B)
	commit

Note -- In schedules 1, 2 and 3, the sum “A + B” is preserved.

Schedule 4

- The following concurrent schedule does not preserve the sum of "$A + B$"

T_1	T_2
read (A)	
$A := A - 50$	
	read (A)
	$temp := A * 0.1$
	$A := A - temp$
	write (A)
	read (B)
write (A)	
read (B)	
$B := B + 50$	
write (B)	
commit	
	$B := B + temp$
	write (B)
	commit

Serializability

- **Basic Assumption** – Each transaction preserves database consistency.
- Thus, serial execution of a set of transactions preserves database consistency.
- A (possibly concurrent) schedule is serializable if it is equivalent to a serial schedule. Different forms of schedule equivalence give rise to the notions of:
  1. **conflict serializability**
  2. **view serializability**

Simplified view of transactions

- We ignore operations other than **read** and **write** instructions
- We assume that transactions may perform arbitrary computations on data in local buffers in between reads and writes.
- Our simplified schedules consist of only **read** and **write** instructions.

Conflicting Instructions

- Let I_i and I_j be two Instructions of transactions T_i and T_j respectively. Instructions I_i and I_j **conflict** if and only if there exists some item Q accessed by both I_i and I_j, and at least one of these instructions wrote Q.

  1. I_i = **read**(Q), I_j = **read**(Q). I_i and I_j don't conflict.
  2. I_i = **read**(Q), I_j = **write**(Q). They conflict.
  3. I_i = **write**(Q), I_j = **read**(Q). They conflict
  4. I_i = **write**(Q), I_j = **write**(Q). They conflict

- Intuitively, a conflict between I_i and I_j forces a (logical) temporal order between them.
  - If I_i and I_j are consecutive in a schedule and they do not conflict, their results would remain the same even if they had been interchanged in the schedule.

Conflict Serializability

- If a schedule *S* can be transformed into a schedule *S´* by a series of swaps of non-conflicting instructions, we say that *S* and *S´* are **conflict equivalent**.
- We say that a schedule *S* is **conflict serializable** if it is conflict equivalent to a serial schedule

Conflict Serializability (Cont.)

- Schedule 3 can be transformed into Schedule 6 -- a serial schedule where T_2 follows T_1, by a series of swaps of non-conflicting instructions. Therefore, Schedule 3 is conflict serializable.

T_1	T_2
read (A)	
write (A)	
	read (A)
	write (A)
read (B)	
write (B)	
	read (B)
	write (B)

Schedule 3

T_1	T_2
read (A)	
write (A)	
read (B)	
write (B)	
	read (A)
	write (A)
	read (B)
	write (B)

Schedule 6

Conflict Serializability (Cont.)

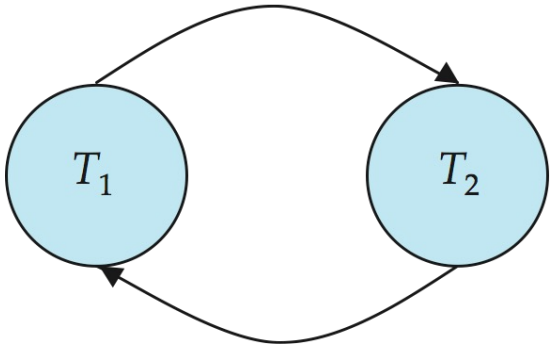
- Example of a schedule that is not conflict serializable:

T_3	T_4
read (Q)	
	write (Q)
write (Q)	

- We are unable to swap instructions in the above schedule to obtain either the serial schedule < T_3, T_4 >, or the serial schedule < T_4, T_3 >.

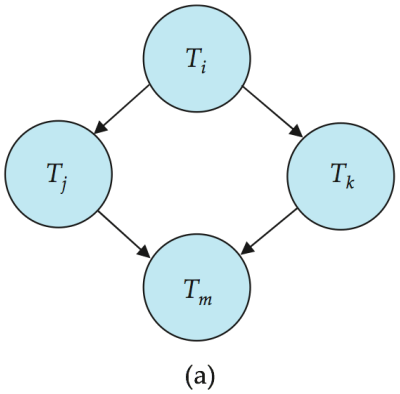
Precedence Graph

- Consider some schedule of a set of transactions T_1, T_2, ..., T_n
- **Precedence graph** — a direct graph where the vertices are the transactions (names).
- We draw an arc from T_i to T_j if the two transaction conflict, and T_i accessed the data item on which the conflict arose earlier.
- We may label the arc by the item that was accessed.
- **Example**

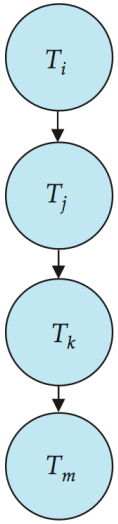


Testing for Conflict Serializability

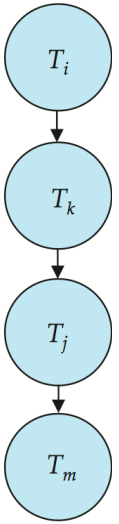
- A schedule is conflict serializable if and only if its precedence graph is acyclic.
- Cycle-detection algorithms exist which take order n^2 time, where n is the number of vertices in the graph.
  - (Better algorithms take order $n + e$ where e is the number of edges.)
- If precedence graph is acyclic, the serializability order can be obtained by a *topological sorting* of the graph.
  - That is, a linear order consistent with the partial order of the graph.
  - For example, a serializability order for the schedule (a) would be one of either (b) or (c)



(a)



(b)



(c)

Recoverable Schedules

- **Recoverable schedule** — if a transaction T_j reads a data item previously written by a transaction T_i, then the commit operation of T_i **must** appear before the commit operation of T_j.
- The following schedule is not recoverable if T_9 commits immediately after the read(A) operation.

T_8	T_9
read (A)	
write (A)	
	read (A)
	commit
read (B)	

- If T_8 should abort, T_9 would have read (and possibly shown to the user) an inconsistent database state. Hence, database must ensure that schedules are recoverable.

Cascading Rollbacks

- **Cascading rollback** – a single transaction failure leads to a series of transaction rollbacks. Consider the following schedule where none of the transactions has yet committed (so the schedule is recoverable)

T_{10}	T_{11}	T_{12}
read (A)		
read (B)		
write (A)		
	read (A)	
	write (A)	
		read (A)
abort		

If T_{10} fails, T_{11} and T_{12} must also be rolled back.

- Can lead to the undoing of a significant amount of work

Cascadeless Schedules

- **Cascadeless schedules** — for each pair of transactions T_i and T_j such that T_j reads a data item previously written by T_i, the commit operation of T_i appears before the read operation of T_j.
- Every cascadeless schedule is also recoverable
- It is desirable to restrict the schedules to those that are cascadeless
- Example of a schedule that is NOT cascadeless

T_{10}	T_{11}	T_{12}
read (A)		
read (B)		
write (A)		
	read (A)	
	write (A)	
		read (A)
abort		

Concurrency Control

- A database must provide a mechanism that will ensure that all possible schedules are both:
  - Conflict serializable.
  - Recoverable and preferably cascadeless
- A policy in which only one transaction can execute at a time generates serial schedules, but provides a poor degree of concurrency
- Concurrency-control schemes tradeoff between the amount of concurrency they allow and the amount of overhead that they incur
- Testing a schedule for serializability *after* it has executed is a little too late!
  - Tests for serializability help us understand why a concurrency control protocol is correct
- **Goal** – to develop concurrency control protocols that will assure serializability.

Weak Levels of Consistency

- Some applications are willing to live with weak levels of consistency, allowing schedules that are not serializable
  - E.g., a read-only transaction that wants to get an approximate total balance of all accounts
  - E.g., database statistics computed for query optimization can be approximate (why?)
  - Such transactions need not be serializable with respect to other transactions
- Tradeoff accuracy for performance

Levels of Consistency in SQL-92

- **Serializable** — default
- **Repeatable read** — only committed records to be read, repeated reads of same record must return same value. However, a transaction may not be serializable – it may find some records inserted by a transaction but not find others.
- **Read committed** — only committed records can be read, but successive reads of record may return different (but committed) values.
- **Read uncommitted** — even uncommitted records may be read.

- Lower degrees of consistency useful for gathering approximate information about the database
- Warning: some database systems do not ensure serializable schedules by default
  - E.g., Oracle and PostgreSQL by default support a level of consistency called snapshot isolation (not part of the SQL standard)

Transaction Definition in SQL

- Data manipulation language must include a construct for specifying the set of actions that comprise a transaction.
- In SQL, a transaction begins implicitly.
- A transaction in SQL ends by:
  - **Commit work** commits current transaction and begins a new one.
  - **Rollback work** causes current transaction to abort.
- In almost all database systems, by default, every SQL statement also commits implicitly if it executes successfully
  - Implicit commit can be turned off by a database directive
    - E.g. in JDBC, connection.setAutoCommit(false);

Other Notions of Serializability

View Serializability

- Let *S* and *S´* be two schedules with the same set of transactions. *S* and *S´* are **view equivalent** if the following three conditions are met, for each data item *Q,*
  1. If in schedule S, transaction T_i reads the initial value of *Q*, then in schedule *S'* also transaction T_i must read the initial value of *Q.*
  2. If in schedule S transaction T_i executes **read**(*Q*), and that value was produced by transaction T_j (if any), then in schedule *S'* also transaction T_i must read the value of *Q* that was produced by the same **write**(Q) operation of transaction T_j .
  3. The transaction (if any) that performs the final **write**(*Q*) operation in schedule *S* must also perform the final **write**(*Q*) operation in schedule *S'.*
- As can be seen, view equivalence is also based purely on **reads** and **writes** alone.

View Serializability (Cont.)

- A schedule S is **view serializable** if it is view equivalent to a serial schedule.
- Every conflict serializable schedule is also view serializable.
- Below is a schedule which is view-serializable but *not* conflict serializable.

T_{27}	T_{28}	T_{29}
read (Q)		
	write (Q)	
write (Q)		
		write (Q)

- What serial schedule is above equivalent to?
- Every view serializable schedule that is not conflict serializable has **blind writes**.

Test for View Serializability

- The precedence graph test for conflict serializability cannot be used directly to test for view serializability.
  - Extension to test for view serializability has cost exponential in the size of the precedence graph.
- The problem of checking if a schedule is view serializable falls in the class of *NP*-complete problems.
  - Thus, existence of an efficient algorithm is *extremely* unlikely.
- However ,practical algorithms that just check some **sufficient conditions** for view serializability can still be used.

More Complex Notions of Serializability

- The schedule below produces the same outcome as the serial schedule $< T_1, T_5 >$, yet is not conflict equivalent or view equivalent to it.

T_1	T_5
read (A)	
$A := A - 50$	
write (A)	
	read (B)
	$B := B - 10$
	write (B)
read (B)	
$B := B + 50$	
write (B)	
	read (A)
	$A := A + 10$
	write (A)

- If we start with A = 1000 and B = 2000, the final result is 960 and 2040
- Determining such equivalence requires analysis of operations other than read and write.

Concurrency Control

Lock-Based Protocols

- A lock is a mechanism to control concurrent access to a data item
- Data items can be locked in two modes :
  1. *exclusive (X) mode*. Data item can be both read as well as written. X-lock is requested using **lock-X** instruction.
  2. *shared (S) mode*. Data item can only be read. S-lock is requested using **lock-S** instruction.
- Lock requests are made to the concurrency-control manager by the programmer. Transaction can proceed only after request is granted.

Lock-Based Protocols (Cont.)

- **Lock-compatibility matrix**

	S	X
S	true	false
X	false	false

- A transaction may be granted a lock on an item if the requested lock is compatible with locks already held on the item by other transactions
- Any number of transactions can hold shared locks on an item,
  - But if any transaction holds an exclusive on the item no other transaction may hold any lock on the item.
- If a lock cannot be granted, the requesting transaction is made to wait till all incompatible locks held by other transactions have been released. The lock is then granted.

Lock-Based Protocols (Cont.)

- Example of a transaction performing locking:

  T_2: **lock-S**(*A*);
  read (*A*);
  unlock(*A*);
  lock-S(*B*);
  read (*B*);
  unlock(*B*);
  display(*A*+*B*)

- Locking as above is not sufficient to guarantee serializability — if *A* and *B* get updated in-between the read of *A* and *B*, the displayed sum would be wrong.
- A **locking protocol** is a set of rules followed by all transactions while requesting and releasing locks. Locking protocols restrict the set of possible schedules.

The Two-Phase Locking Protocol

- This protocol ensures conflict-serializable schedules.
- Phase 1: Growing Phase
  - Transaction may obtain locks
  - Transaction may not release locks
- Phase 2: Shrinking Phase
  - Transaction may release locks
  - Transaction may not obtain locks
- The protocol assures serializability. It can be proved that the transactions can be serialized in the order of their **lock points** (i.e., the point where a transaction acquired its final lock).

The Two-Phase Locking Protocol (Cont.)

- There can be conflict serializable schedules that cannot be obtained if two-phase locking is used.
- However, in the absence of extra information (e.g., ordering of access to data), two-phase locking is needed for conflict serializability in the following sense:
  - Given a transaction T_i that does not follow two-phase locking, we can find a transaction T_j that uses two-phase locking, and a schedule for T_i and T_j that is not conflict serializable.

Lock Conversions

- Two-phase locking with lock conversions:
  - First Phase:
    - can acquire a lock-S on item
    - can acquire a lock-X on item
    - can convert a lock-S to a lock-X (upgrade)
  - Second Phase:
    - can release a lock-S
    - can release a lock-X
    - can convert a lock-X to a lock-S (downgrade)
- This protocol assures serializability. But still relies on the programmer to insert the various locking instructions.

Automatic Acquisition of Locks

- A transaction T_i issues the standard read/write instruction, without explicit locking calls.
- The operation **read**(*D*) is processed as:

```
if Ti has a lock on D
  then
      read(D)
  else begin
      if necessary wait until no other
         transaction has a lock-X on D
      grant Ti a lock-S on D;
      read(D)
    end
```

Automatic Acquisition of Locks (Cont.)

- **write**(*D*) is processed as:

```
if Ti has a lock-X on D
  then
    write(D)
  else begin
    if necessary wait until no other transaction has any lock on D,
    if Ti has a lock-S on D
      then
        upgrade lock on D to lock-X
      else
        grant Ti a lock-X on D
    write(D)
  end;
```

- All locks are released after commit or abort

Deadlocks

- Consider the partial schedule

T_3	T_4
lock-x (B)	
read (B)	
$B := B - 50$	
write (B)	
	lock-s (A)
	read (A)
	lock-s (B)
lock-x (A)	

- Neither T_3 nor T_4 can make progress — executing **lock-S**(*B*) causes T_4 to wait for T_3 to release its lock on *B*, while executing **lock-X**(*A*) causes T_3 to wait for T_4 to release its lock on *A*.
- Such a situation is called a **deadlock**.
  - To handle a deadlock one of T_3 or T_4 must be rolled back and its locks released.

Deadlocks (Cont.)

- Two-phase locking *does not* ensure freedom from deadlocks.
- In addition to deadlocks, there is a possibility of **starvation.**
- **Starvation** occurs if the concurrency control manager is badly designed. For example:
  - A transaction may be waiting for an X-lock on an item, while a sequence of other transactions request and are granted an S-lock on the same item.
  - The same transaction is repeatedly rolled back due to deadlocks.
- Concurrency control manager can be designed to prevent starvation.

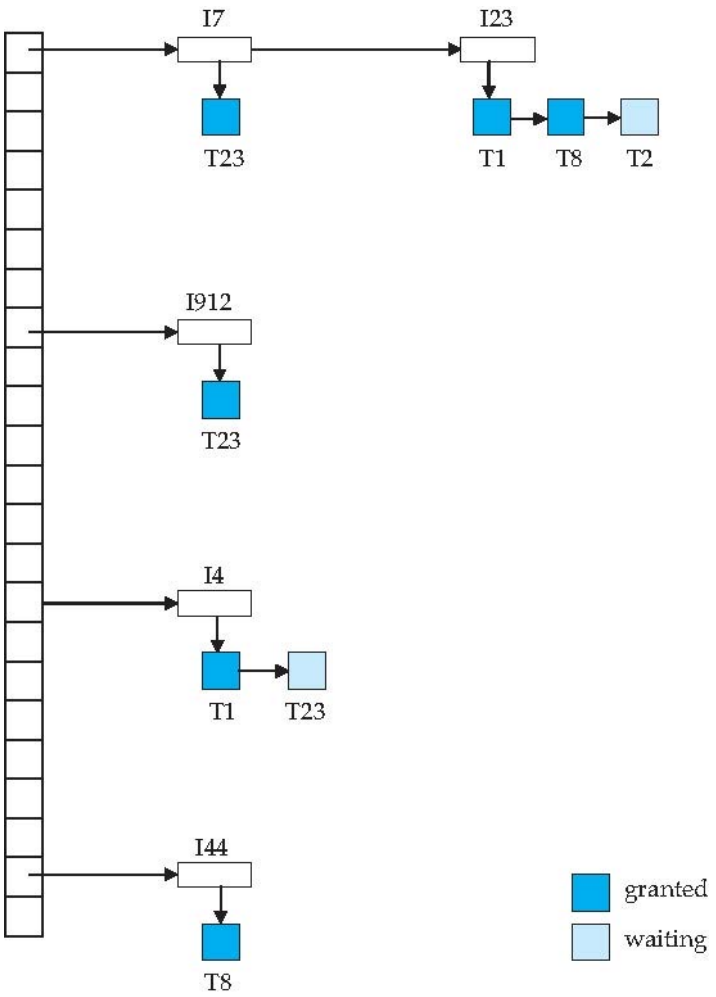
Deadlocks (Cont.)

- The potential for deadlock exists in most locking protocols. Deadlocks are a necessary evil.
- When a deadlock occurs there is a possibility of cascading roll-backs.
- Cascading roll-back is possible under two-phase locking. To avoid this, follow a modified protocol called **strict two-phase locking** -- a transaction must hold all its exclusive locks till it commits/aborts.
- **Rigorous two-phase locking** is even stricter. Here, *all* locks are held till commit/abort. In this protocol transactions can be serialized in the order in which they commit.

Implementation of Locking

- A **lock manager** can be implemented as a separate process to which transactions send lock and unlock requests
- The lock manager replies to a lock request by sending a lock grant messages (or a message asking the transaction to roll back, in case of a deadlock)
- The requesting transaction waits until its request is answered
- The lock manager maintains a data-structure called a **lock table** to record granted locks and pending requests
- The lock table is usually implemented as an in-memory hash table indexed on the name of the data item being locked

Lock Table



- Dark blue rectangles indicate granted locks; light blue indicate waiting requests
- Lock table also records the type of lock granted or requested
- New request is added to the end of the queue of requests for the data item, and granted if it is compatible with all earlier locks
- Unlock requests result in the request being deleted, and later requests are checked to see if they can now be granted
- If transaction aborts, all waiting or granted requests of the transaction are deleted
  - lock manager may keep a list of locks held by each transaction, to implement this efficiently

Deadlock Handling

- System is deadlocked if there is a set of transactions such that every transaction in the set is waiting for another transaction in the set.
- ***Deadlock prevention*** protocols ensure that the system will *never* enter into a deadlock state. Some prevention strategies :
  - Require that each transaction locks all its data items before it begins execution (predeclaration).
  - Impose partial ordering of all data items and require that a transaction can lock data items only in the order specified by the partial order.

More Deadlock Prevention Strategies

- Following schemes use transaction timestamps for the sake of deadlock prevention alone.
- **wait-die** scheme — non-preemptive
  - older transaction may wait for younger one to release data item. (older means smaller timestamp) Younger transactions never Younger transactions never wait for older ones; they are rolled back instead.
  - a transaction may die several times before acquiring needed data item
- **wound-wait** scheme — preemptive
  - older transaction *wounds* (forces rollback) of younger transaction instead of waiting for it. Younger transactions may wait for older ones.
  - may be fewer rollbacks than *wait-die* scheme.

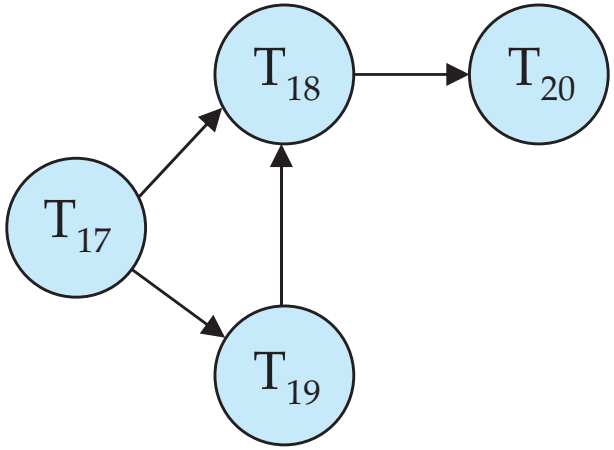
Deadlock prevention (Cont.)

- Both in *wait-die* and in *wound-wait* schemes, a rolled back transactions is restarted with its original timestamp. Older transactions thus have precedence over newer ones, and starvation is hence avoided.
- **Timeout-Based Schemes:**
  - a transaction waits for a lock only for a specified amount of time. If the lock has not been granted within that time, the transaction is rolled back and restarted,
  - Thus, deadlocks are not possible
  - simple to implement; but starvation is possible. Also difficult to determine good value of the timeout interval.

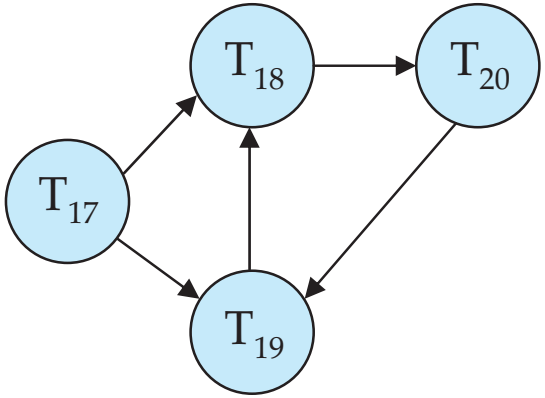
Deadlock Detection

- Deadlocks can be described as a *wait-for graph*, which consists of a pair $G = (V,E)$,
  - V is a set of vertices (all the transactions in the system)
  - E is a set of edges; each element is an ordered pair $T_i \rightarrow T_j$.
- If $T_i \rightarrow T_j$ is in E, then there is a directed edge from T_i to T_j, implying that T_i is waiting for T_j to release a data item.
- When T_i requests a data item currently being held by T_j, then the edge $T_i \rightarrow T_j$ is inserted in the wait-for graph. This edge is removed only when T_j is no longer holding a data item needed by T_i.
- The system is in a deadlock state if and only if the wait-for graph has a cycle. Must invoke a deadlock-detection algorithm periodically to look for cycles.

Deadlock Detection (Cont.)



Wait-for graph without a cycle



Wait-for graph with a cycle

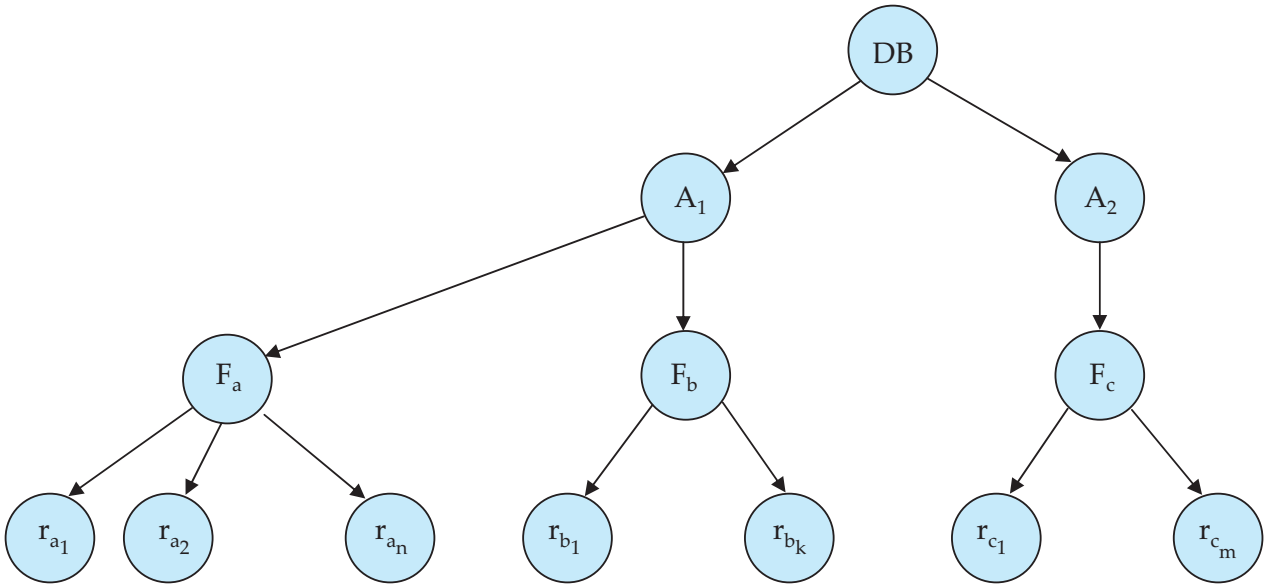
Deadlock Recovery

- When deadlock is detected :
  - Some transaction will have to rolled back (made a victim) to break deadlock. Select that transaction as victim that will incur minimum cost.
  - Rollback -- determine how far to roll back transaction
    - Total rollback: Abort the transaction and then restart it.
    - More effective to roll back transaction only as far as necessary to break deadlock.
  - Starvation happens if same transaction is always chosen as victim. Include the number of rollbacks in the cost factor to avoid starvation

Multiple Granularity

- Allow data items to be of various sizes and define a hierarchy of data granularities, where the small granularities are nested within larger ones
- Can be represented graphically as a tree.
- When a transaction locks a node in the tree *explicitly*, it *implicitly* locks all the node's descendents in the same mode.
- **Granularity of locking** (level in tree where locking is done):
  - fine granularity (lower in tree): high concurrency, high locking overhead
  - coarse granularity (higher in tree): low locking overhead, low concurrency

Example of Granularity Hierarchy



The levels, starting from the coarsest (top) level are

- *database*
- *area*
- *file*
- *record*

Intention Lock Modes

- In addition to S and X lock modes, there are three additional lock modes with multiple granularity:
  - ***intention-shared*** (IS): indicates explicit locking at a lower level of the tree but only with shared locks.
  - ***intention-exclusive*** (IX): indicates explicit locking at a lower level with exclusive or shared locks
  - ***shared and intention-exclusive*** (SIX): the subtree rooted by that node is locked explicitly in shared mode and explicit locking is being done at a lower level with exclusive-mode locks.
- intention locks allow a higher level node to be locked in S or X mode without having to check all descendent nodes.

Compatibility Matrix with Intention Lock Modes

- The compatibility matrix for all lock modes is:

	IS	IX	S	SIX	X
IS	true	true	true	true	false
IX	true	true	false	false	false
S	true	false	true	false	false
SIX	true	false	false	false	false
X	false	false	false	false	false

Multiple Granularity Locking Scheme

- Transaction T_i can lock a node Q, using the following rules:
  1. The lock compatibility matrix must be observed.
  2. The root of the tree must be locked first, and may be locked in any mode.
  3. A node Q can be locked by T_i in S or IS mode only if the parent of Q is currently locked by T_i in either IX or IS mode.
  4. A node Q can be locked by T_i in X, SIX, or IX mode only if the parent of Q is currently locked by T_i in either IX or SIX mode.
  5. T_i can lock a node only if it has not previously unlocked any node (that is, T_i is two-phase).
  6. T_i can unlock a node Q only if none of the children of Q are currently locked by T_i.
- Observe that locks are acquired in root-to-leaf order, whereas they are released in leaf-to-root order.
- **Lock granularity escalation**: in case there are too many locks at a particular level, switch to higher granularity S or X lock

Timestamp-Based Protocols

- Each transaction is issued a timestamp when it enters the system. If an old transaction T_i has time-stamp TS(T_i), a new transaction T_j is assigned time-stamp TS(T_j) such that TS(T_i) <TS(T_j).
- The protocol manages concurrent execution such that the time-stamps determine the serializability order.
- In order to assure such behavior, the protocol maintains for each data Q two timestamp values:
  - **W-timestamp**(Q) is the largest time-stamp of any transaction that executed **write**(Q) successfully.
  - **R-timestamp**(Q) is the largest time-stamp of any transaction that executed **read**(Q) successfully.

Timestamp-Based Protocols (Cont.)

- The timestamp ordering protocol ensures that any conflicting **read** and **write** operations are executed in timestamp order.
- Suppose a transaction T_i issues a **read**(Q)
  1. If TS(T_i) ≤ **W**-timestamp(Q), then T_i needs to read a value of Q that was already overwritten.
     - Hence, the **read** operation is rejected, and T_i is rolled back.
  2. If TS(T_i) ≥ **W**-timestamp(Q), then the **read** operation is executed, and R-timestamp(Q) is set to **max**(R-timestamp(Q), TS(T_i)).

Timestamp-Based Protocols (Cont.)

- Suppose that transaction T_i issues **write**(Q).
  1. If TS(T_i) < R-timestamp(Q), then the value of Q that T_i is producing was needed previously, and the system assumed that that value would never be produced.
     - Hence, the **write** operation is rejected, and T_i is rolled back.
  2. If TS(T_i) < W-timestamp(Q), then T_i is attempting to write an obsolete value of Q.
     - Hence, this **write** operation is rejected, and T_i is rolled back.
  3. Otherwise, the **write** operation is executed, and W-timestamp(Q) is set to TS(T_i).

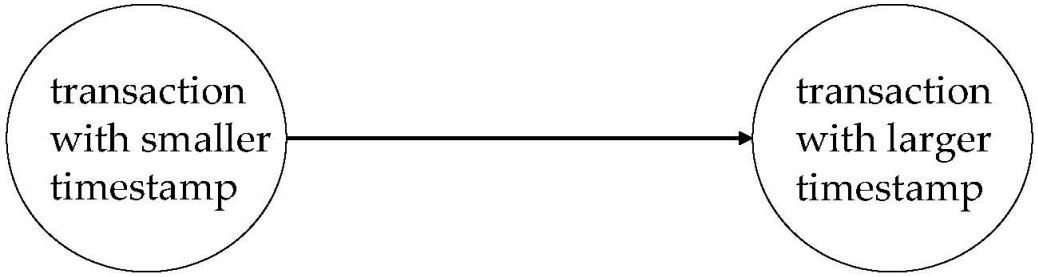
Example Use of the Protocol

A partial schedule for several data items for transactions with timestamps 1, 2, 3, 4, 5

T_1	T_2	T_3	T_4	T_5
				read (X)
	read (Y)			
read (Y)				
		write (Y)		
		write (Z)		
				read (Z)
	read (Z)			
	abort			
read (X)				
			read (W)	
		write (W)		
		abort		
				write (Y)
				write (Z)

Correctness of Timestamp-Ordering Protocol

- The timestamp-ordering protocol guarantees serializability since all the arcs in the precedence graph are of the form:



Thus, there will be no cycles in the precedence graph

- Timestamp protocol ensures freedom from deadlock as no transaction ever waits.
- But the schedule may not be cascade-free, and may not even be recoverable.

Recoverability and Cascade Freedom

- Problem with timestamp-ordering protocol:
  - Suppose T_i aborts, but T_j has read a data item written by T_i
  - Then T_j must abort; if T_j had been allowed to commit earlier, the schedule is not recoverable.
  - Further, any transaction that has read a data item written by T_j must abort
  - This can lead to cascading rollback --- that is, a chain of rollbacks
- Solution 1:
  - A transaction is structured such that its writes are all performed at the end of its processing
  - All writes of a transaction form an atomic action; no transaction may execute while a transaction is being written
  - A transaction that aborts is restarted with a new timestamp
- Solution 2: Limited form of locking: wait for data to be committed before reading it
- Solution 3: Use commit dependencies to ensure recoverability

Thomas' Write Rule

- Modified version of the timestamp-ordering protocol in which obsolete **write** operations may be ignored under certain circumstances.
- When T_i attempts to write data item Q, if TS(T_i) < W-timestamp(Q), then T_i is attempting to write an obsolete value of {Q}.
  - Rather than rolling back T_i as the timestamp ordering protocol would have done, this {**write**} operation can be ignored.
- Otherwise this protocol is the same as the timestamp ordering protocol.
- Thomas' Write Rule allows greater potential concurrency.
  - Allows some view-serializable schedules that are not conflict-serializable.

Validation-Based Protocol

- Execution of transaction T_i is done in three phases.
  1. **Read and execution phase**: Transaction T_i writes only to temporary local variables
  2. **Validation phase**: Transaction T_i performs a "validation test" to determine if local variables can be written without violating serializability.
  3. **Write phase**: If T_i is validated, the updates are applied to the database; otherwise, T_i is rolled back.
- The three phases of concurrently executing transactions can be interleaved, but each transaction must go through the three phases in that order.
  - Assume for simplicity that the validation and write phase occur together, atomically and serially
    - I.e., only one transaction executes validation/write at a time.
- Also called as **optimistic concurrency control** since transaction executes fully in the hope that all will go well during validation

Validation-Based Protocol (Cont.)

- Each transaction T_i has 3 timestamps
  - Start(T_i) : the time when T_i started its execution
  - Validation(T_i): the time when T_i entered its validation phase
  - Finish(T_i) : the time when T_i finished its write phase
- Serializability order is determined by timestamp given at validation time; this is done to increase concurrency.
  - Thus, TS(T_i) is given the value of Validation(T_i).
- This protocol is useful and gives greater degree of concurrency if probability of conflicts is low.
  - because the serializability order is not pre-decided, and
  - relatively few transactions will have to be rolled back.

Validation Test for Transaction T_j

- If for all T_i with TS (T_i) < TS (T_j) either one of the following condition holds:
  - **finish**(T_i) < **start**(T_j)
  - **start**(T_j) < **finish**(T_i) < **validation**(T_j) **and** the set of data items written by T_i does not intersect with the set of data items read by T_j.

  then validation succeeds and T_j can be committed. Otherwise, validation fails and T_j is aborted.
- *Justification*: Either the first condition is satisfied, and there is no overlapped execution, or the second condition is satisfied and
  - the writes of T_j do not affect reads of T_i since they occur after T_i has finished its reads.
  - the writes of T_i do not affect reads of T_j since T_j does not read any item written by T_i.

Schedule Produced by Validation

- Example of schedule produced using validation

T_{25}	T_{26}
read (B)	
	read (B)
	$B := B$ 50
	read (A)
	$A := A + 50$
read (A)	
⟨ *validate* ⟩	
display ($A + B$)	
	⟨ *validate* ⟩
	write (B)
	write (A)

Multiversion Schemes

- Multiversion schemes keep old versions of data item to increase concurrency.
  - Multiversion Timestamp Ordering
  - Multiversion Two-Phase Locking
- Each successful **write** results in the creation of a new version of the data item written.
- Use timestamps to label versions.
- When a **read**(*Q*) operation is issued, select an appropriate version of *Q* based on the timestamp of the transaction, and return the value of the selected version.
- **read**s never have to wait as an appropriate version is returned immediately.

Multiversion Timestamp Ordering

- Each data item Q has a sequence of versions $<Q_1, Q_2, ..., Q_m>$. Each version Q_k contains three data fields:
  - **Content** -- the value of version Q_k.
  - **W-timestamp**(Q_k) -- timestamp of the transaction that created (wrote) version Q_k
  - **R-timestamp**(Q_k) -- largest timestamp of a transaction that successfully read version Q_k
- When a transaction T_i creates a new version Q_k of Q, Q_k's W-timestamp and R-timestamp are initialized to TS(T_i).
- R-timestamp of Q_k is updated whenever a transaction T_j reads Q_k, and TS(T_j) > R-timestamp(Q_k).

Multiversion Timestamp Ordering (Cont)

- Suppose that transaction T_i issues a **read**(Q) or **write**(Q) operation. Let Q_k denote the version of Q whose write timestamp is the largest write timestamp less than or equal to TS(T_i).
  1. If transaction T_i issues a **read**(Q), then the value returned is the content of version Q_k.
  2. If transaction T_i issues a **write**(Q)
     1. if TS(T_i) < R-timestamp(Q_k), then transaction T_i is rolled back.
     2. if TS(T_i) = W-timestamp(Q_k), the contents of Q_k are overwritten
     3. else a new version of Q is created.
- Observe that
  - Reads always succeed
  - A write by T_i is rejected if some other transaction T_j that (in the serialization order defined by the timestamp values) should read T_i's write, has already read a version created by a transaction older than T_i.
- Protocol guarantees serializability

Multiversion Two-Phase Locking

- Differentiates between read-only transactions and update transactions
- *Update transactions* acquire read and write locks, and hold all locks up to the end of the transaction. That is, update transactions follow rigorous two-phase locking.
  - Each successful **write** results in the creation of a new version of the data item written.
  - Each version of a data item has a single timestamp whose value is obtained from a counter **ts-counter** that is incremented during commit processing.
- *Read-only transactions* are assigned a timestamp by reading the current value of **ts-counter** before they start execution; they follow the multiversion timestamp-ordering protocol for performing reads.

Multiversion Two-Phase Locking (Cont.)

- When an update transaction wants to read a data item:
  - it obtains a shared lock on it, and reads the latest version.
- When it wants to write an item
  - it obtains X lock on; it then creates a new version of the item and sets this version's timestamp to ∞.
- When update transaction T_i completes, commit processing occurs:
  - T_i sets timestamp on the versions it has created to **ts-counter** + 1
  - T_i increments **ts-counter** by 1
- Read-only transactions that start after T_i increments **ts-counter** will see the values updated by T_i.
- Read-only transactions that start before T_i increments the **ts-counter** will see the value before the updates by T_i.
- Only serializable schedules are produced.

MVCC: Implementation Issues

- Creation of multiple versions increases storage overhead
  - Extra tuples
  - Extra space in each tuple for storing version information
- Versions can, however, be garbage collected
  - E.g. if Q has two versions Q5 and Q9, and the oldest active transaction has timestamp > 9, than Q5 will never be required again

Snapshot Isolation

- Motivation: Decision support queries that read large amounts of data have concurrency conflicts with OLTP transactions that update a few rows
  - Poor performance results
- Solution 1: Give logical "snapshot" of database state to read only transactions, read-write transactions use normal locking
  - Multiversion 2-phase locking
  - Works well, but how does system know a transaction is read only?
- Solution 2: Give snapshot of database state to every transaction, updates alone use 2-phase locking to guard against concurrent updates
  - Problem: variety of anomalies such as lost update can result
  - Partial solution: snapshot isolation level (next slide)
    - Proposed by Berenson et al, SIGMOD 1995
    - Variants implemented in many database systems
      - E.g. Oracle, PostgreSQL, SQL Server 2005

Snapshot Isolation

- A transaction T1 executing with Snapshot Isolation
  - takes snapshot of committed data at start
  - always reads/modifies data in its own snapshot
  - updates of concurrent transactions are not visible to T1
  - writes of T1 complete when it commits
  - **First-committer-wins rule**:
    - Commits only if no other concurrent transaction has already written data that T1 intends to write.

T1	T2	T3
W(Y := 1) Commit		
	Start R(X) → 0 R(Y)→ 1	
		W(X:=2) W(Z:=3) Commit
	R(Z) → 0 R(Y) → 1 W(X:=3) Commit-Req Abort	

Concurrent updates not visible
Own updates are visible
Not first-committer of X
Serialization error, T2 is rolled back

Snapshot Read

- Concurrent updates invisible to snapshot read

$X_0 = 100,\ Y_0 = 0$

T_1 deposits 50 in Y	T_2 withdraws 50 from X
$r_1(X_0, 100)$	
$r_1(Y_0, 0)$	
	$r_2(Y_0, 0)$
	$r_2(X_0, 100)$
	$w_2(X_2, 50)$
$w_1(Y_1, 50)$	
$r_1(X_0, 100)$ (update by T_2 not seen)	
$r_1(Y_1, 50)$ (can see its own updates)	
	$r_2(Y_0, 0)$ (update by T_1 not seen)

$X_2 = 50,\ Y_1 = 50$

Snapshot Write: First Committer Wins

$X_0 = 100$

T_1 deposits 50 in X	T_2 withdraws 50 from X
$r_1(X_0, 100)$	
	$r_2(X_0, 100)$
	$w_2(X_2, 50)$
$w_1(X_1, 150)$	
$commit_1$	
	$commit_2$ (Serialization Error T_2 is rolled back)

$X_1 = 150$

- Variant: "**First-updater-wins**"
  - Check for concurrent updates when write occurs by locking item
    - But lock should be held till all concurrent transactions have finished
  - (Oracle uses this plus some extra features)
  - Differs only in when abort occurs, otherwise equivalent

Benefits of Snapshot Isolation

- Reading is *never* blocked,
  - and also doesn't block other txns activities
- Performance similar to Read Committed
- Avoids the usual anomalies
  - No dirty read
  - No lost update
  - No non-repeatable read
  - Predicate based selects are repeatable (no phantoms)
- Problems with SI
  - SI does not always give serializable executions
    - Serializable: among two concurrent txns, one sees the effects of the other
    - In SI: neither sees the effects of the other
  - Result: Integrity constraints can be violated

Snapshot Isolation

- E.g. of problem with SI
  - T1: x:=y
  - T2: y:= x
  - Initially x = 3 and y = 17
    - Serial execution: x = ??, y = ??
    - if both transactions start at the same time, with snapshot isolation: x = ?? , y = ??
- Called **skew write**
- Skew also occurs with inserts
  - E.g:
    - Find max order number among all orders
    - Create a new order with order number = previous max + 1

Snapshot Isolation Anomalies

- SI breaks serializability when txns modify *different* items, each based on a previous state of the item the other modified
  - Not very common in practice
    - E.g., the TPC-C benchmark runs correctly under SI
    - when txns conflict due to modifying different data, there is usually also a shared item they both modify too (like a total quantity) so SI will abort one of them
  - But does occur
    - Application developers should be careful about write skew
- SI can also cause a read-only transaction anomaly, where read-only transaction may see an inconsistent state even if updaters are serializable
  - We omit details
- Using snapshots to verify primary/foreign key integrity can lead to inconsistency
  - Integrity constraint checking usually done outside of snapshot

SI In Oracle and PostgreSQL

- **Warning**: SI used when isolation level is set to serializable, by Oracle, and PostgreSQL versions prior to 9.1
  - PostgreSQL's implementation of SI (versions prior to 9.1) described in Section 26.4.1.3
  - Oracle implements "first updater wins" rule (variant of "first committer wins")
    - concurrent writer check is done at time of write, not at commit time
    - Allows transactions to be rolled back earlier
    - Oracle and PostgreSQL < 9.1 do not support true serializable execution
  - PostgreSQL 9.1 introduced new protocol called "Serializable Snapshot Isolation" (SSI)
    - Which guarantees true serializabilty including handling predicate reads (coming up)

SI In Oracle and PostgreSQL

- Can sidestep SI for specific queries by using **select .. for update** in Oracle and PostgreSQL
  - E.g.,
    1. **select max**(orderno) **from** orders **for update**
    2. read value into local variable maxorder
    3. insert into orders (maxorder+1, …)
  - Select for update (SFU) treats all data read by the query as if it were also updated, preventing concurrent updates
  - Does not always ensure serializability since phantom phenomena can occur (coming up)
- In PostgreSQL versions < 9.1, SFU locks the data item, but releases locks when the transaction completes, even if other concurrent transactions are active
  - Not quite same as SFU in Oracle, which keeps locks until all
  - concurrent transactions have completed

Insert and Delete Operations

- If two-phase locking is used :
  - A **delete** operation may be performed only if the transaction deleting the tuple has an exclusive lock on the tuple to be deleted.
  - A transaction that inserts a new tuple into the database is given an X-mode lock on the tuple
- Insertions and deletions can lead to the **phantom phenomenon**.
  - A transaction that scans a relation
    - (e.g., find sum of balances of all accounts in Perryridge)

    and a transaction that inserts a tuple in the relation
    - (e.g., insert a new account at Perryridge)

    (conceptually) conflict in spite of not accessing any tuple in common.
  - If only tuple locks are used, non-serializable schedules can result
    - E.g. the scan transaction does not see the new account, but reads some other tuple written by the update transaction

Insert and Delete Operations (Cont.)

- The transaction scanning the relation is reading information that indicates what tuples the relation contains, while a transaction inserting a tuple updates the same information.
  - The conflict should be detected, e.g. by locking the information.
- One solution:
  - Associate a data item with the relation, to represent the information about what tuples the relation contains.
  - Transactions scanning the relation acquire a shared lock in the data item,
  - Transactions inserting or deleting a tuple acquire an exclusive lock on the data item. (Note: locks on the data item do not conflict with locks on individual tuples.)
- Above protocol provides very low concurrency for insertions/deletions.
- Index locking protocols provide higher concurrency while preventing the phantom phenomenon, by requiring locks on certain index buckets.

Index Locking Protocol

- Index locking protocol:
  - Every relation must have at least one index.
  - A transaction can access tuples only after finding them through one or more indices on the relation
  - A transaction T_i that performs a lookup must lock all the index leaf nodes that it accesses, in S-mode
    - Even if the leaf node does not contain any tuple satisfying the index lookup (e.g. for a range query, no tuple in a leaf is in the range)
  - A transaction T_i that inserts, updates or deletes a tuple t_i in a relation r
    - must update all indices to r
    - must obtain exclusive locks on all index leaf nodes affected by the insert/update/delete
  - The rules of the two-phase locking protocol must be observed
- Guarantees that phantom phenomenon won't occur

Next-Key Locking

- Index-locking protocol to prevent phantoms required locking entire leaf
  - Can result in poor concurrency if there are many inserts
- Alternative: for an index lookup
  - Lock all values that satisfy index lookup (match lookup value, or fall in lookup range)
  - Also lock next key value in index
  - Lock mode: S for lookups, X for insert/delete/update
- Ensures that range queries will conflict with inserts/deletes/updates
  - Regardless of which happens first, as long as both are concurrent

Concurrency in Index Structures

- Indices are unlike other database items in that their only job is to help in accessing data.
- Index-structures are typically accessed very often, much more than other database items.
  - Treating index-structures like other database items, e.g. by 2-phase locking of index nodes can lead to low concurrency.
- There are several index concurrency protocols where locks on internal nodes are released early, and not in a two-phase fashion.
  - It is acceptable to have nonserializable concurrent access to an index as long as the accuracy of the index is maintained.
    - In particular, the exact values read in an internal node of a B^+-tree are irrelevant so long as we land up in the correct leaf node.

Concurrency in Index Structures (Cont.)

- Example of index concurrency protocol:
- Use **crabbing** instead of two-phase locking on the nodes of the B$^+$-tree, as follows. During search/insertion/deletion:
  - First lock the root node in shared mode.
  - After locking all required children of a node in shared mode, release the lock on the node.
  - During insertion/deletion, upgrade leaf node locks to exclusive mode.
  - When splitting or coalescing requires changes to a parent, lock the parent in exclusive mode.
- Above protocol can cause excessive deadlocks
  - Searches coming down the tree deadlock with updates going up the tree
  - Can abort and restart search, without affecting transaction
- Better protocols are available; see Section 16.9 for one such protocol, the B-link tree protocol
  - Intuition: release lock on parent before acquiring lock on child
    - And deal with changes that may have happened between lock release and acquire

Weak Levels of Consistency

- **Degree-two consistency:** differs from two-phase locking in that S-locks may be released at any time, and locks may be acquired at any time
  - X-locks must be held till end of transaction
  - Serializability is not guaranteed, programmer must ensure that no erroneous database state will occur]
- **Cursor stability**:
  - For reads, each tuple is locked, read, and lock is immediately released
  - X-locks are held till end of transaction
  - Special case of degree-two consistency

Weak Levels of Consistency in SQL

- SQL allows non-serializable executions
  - **Serializable:** is the default
  - **Repeatable read**: allows only committed records to be read, and repeating a read should return the same value (so read locks should be retained)
    - However, the phantom phenomenon need not be prevented
      - T1 may see some records inserted by T2, but may not see others inserted by T2
  - **Read committed**: same as degree two consistency, but most systems implement it as cursor-stability
  - **Read uncommitted**: allows even uncommitted data to be read
- In many database systems, read committed is the default consistency level
  - has to be explicitly changed to serializable when required
    - **set isolation level serializable**

Transactions across User Interaction

- Many applications need transaction support across user interactions
  - Can't use locking
  - Don't want to reserve database connection per user
- Application level concurrency control
  - Each tuple has a version number
  - Transaction notes version number when reading tuple
    - **select** r.balance, r.version **into** :A, :version
      from r **where** acctId =23
  - When writing tuple, check that current version number is same as the version when tuple was read
    - **update** r **set** r.balance = r.balance + :deposit
      where acctId = 23 **and** r.version = :version
- Equivalent to **optimistic concurrency control without validating read set**
- Used internally in Hibernate ORM system, and manually in many applications
- Version numbering can also be used to support first committer wins check of snapshot isolation
  - Unlike SI, reads are not guaranteed to be from a single snapshot

Deadlocks

- Consider the following two transactions:

  T_1: write (X) T_2: write(Y)

  write(Y) write(X)

- Schedule with deadlock

T_1	T_2
lock-X on A	
write (A)	
	lock-X on B
	write (B)
	wait for **lock-X** on A
wait for **lock-X** on B	

Recovery System

- Failure Classification
- Storage Structure
- Recovery and Atomicity
- Log-Based Recovery
- Remote Backup Systems

Failure Classification

- **Transaction failure** :
  - **Logical errors**: transaction cannot complete due to some internal error condition
  - **System errors**: the database system must terminate an active transaction due to an error condition (e.g., deadlock)
- **System crash**: a power failure or other hardware or software failure causes the system to crash.
  - **Fail-stop assumption**: non-volatile storage contents are assumed to not be corrupted by system crash
    - Database systems have numerous integrity checks to prevent corruption of disk data
- **Disk failure**: a head crash or similar disk failure destroys all or part of disk storage
  - Destruction is assumed to be detectable: disk drives use checksums to detect failures

Recovery Algorithms

- Consider transaction T_i that transfers \$50 from account *A* to account *B*
  - Two updates: subtract 50 from A and add 50 to B
- Transaction T_i requires updates to A and B to be output to the database.
  - A failure may occur after one of these modifications have been made but before both of them are made.
  - Modifying the database without ensuring that the transaction will commit may leave the database in an inconsistent state
  - Not modifying the database may result in lost updates if failure occurs just after transaction commits
- Recovery algorithms have two parts
  1. Actions taken during normal transaction processing to ensure enough information exists to recover from failures
  2. Actions taken after a failure to recover the database contents to a state that ensures atomicity, consistency and durability

Storage Structure

- **Volatile storage**:
  - does not survive system crashes
  - examples: main memory, cache memory
- **Nonvolatile storage**:
  - survives system crashes
  - examples: disk, tape, flash memory,
    non-volatile (battery backed up) RAM
  - but may still fail, losing data
- **Stable storage**:
  - a mythical form of storage that survives all failures
  - approximated by maintaining multiple copies on distinct nonvolatile media
  - See book for more details on how to implement stable storage

Stable-Storage Implementation

- Maintain multiple copies of each block on separate disks
  - copies can be at remote sites to protect against disasters such as fire or flooding.
- Failure during data transfer can still result in inconsistent copies: Block transfer can result in
  - Successful completion
  - Partial failure: destination block has incorrect information
  - Total failure: destination block was never updated
- Protecting storage media from failure during data transfer (one solution):
  - Execute output operation as follows (assuming two copies of each block):
    1. Write the information onto the first physical block.
    2. When the first write successfully completes, write the same information onto the second physical block.
    3. The output is completed only after the second write successfully completes.

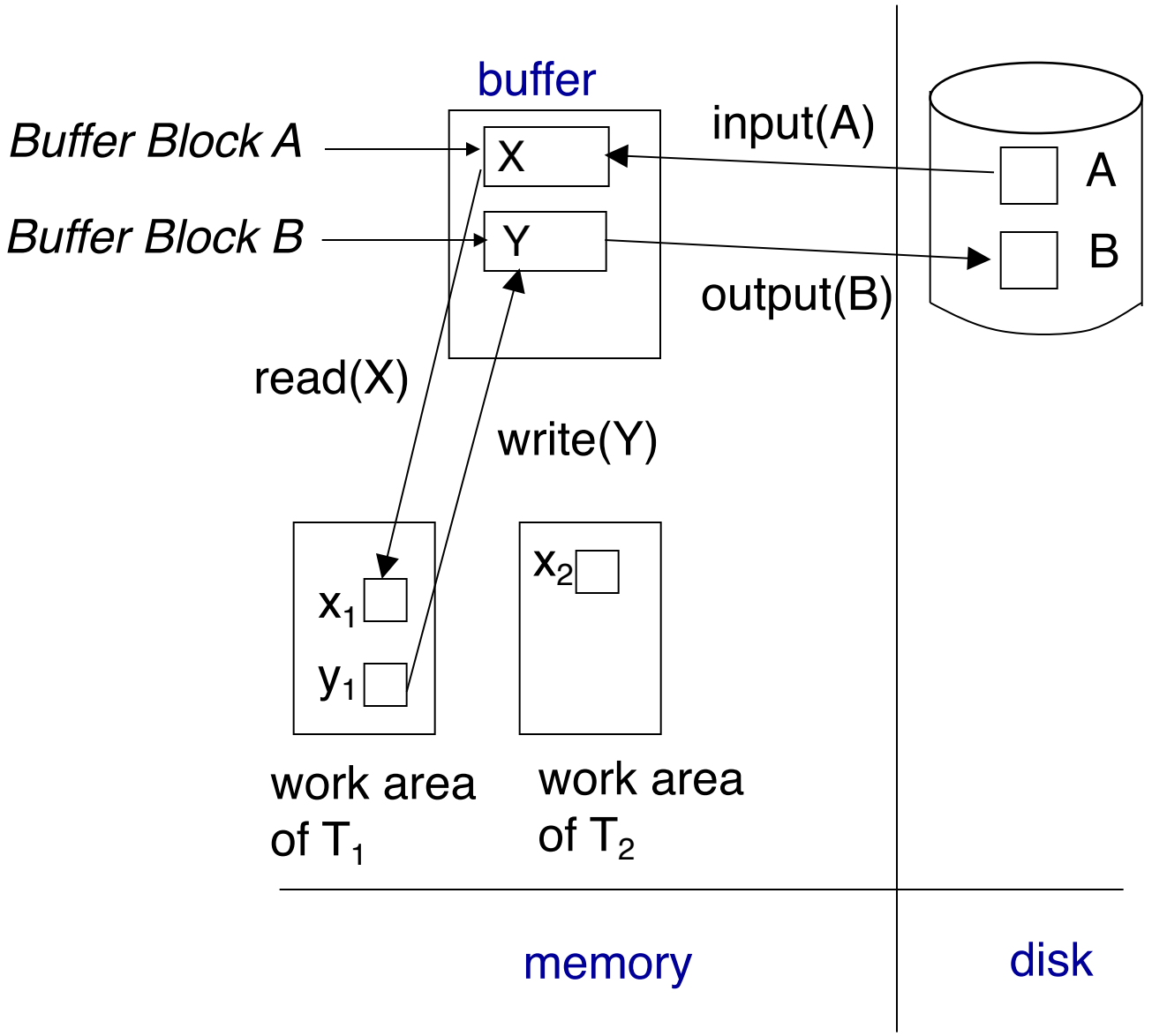
Stable-Storage Implementation (Cont.)

- Protecting storage media from failure during data transfer (cont.):
- Copies of a block may differ due to failure during output operation. To recover from failure:
  1. First find inconsistent blocks:
     1. *Expensive solution*: Compare the two copies of every disk block.
     2. *Better solution*:
        - Record in-progress disk writes on non-volatile storage (Non-volatile RAM or special area of disk).
        - Use this information during recovery to find blocks that may be inconsistent, and only compare copies of these.
        - Used in hardware RAID systems
  2. If either copy of an inconsistent block is detected to have an error (bad checksum), overwrite it by the other copy. If both have no error, but are different, overwrite the second block by the first block.

Data Access

- **Physical blocks** are those blocks residing on the disk.
- **Buffer blocks** are the blocks residing temporarily in main memory.
- Block movements between disk and main memory are initiated through the following two operations:
  - **input**(*B*) transfers the physical block *B* to main memory.
  - **output**(*B*) transfers the buffer block *B* to the disk, and replaces the appropriate physical block there.
- We assume, for simplicity, that each data item fits in, and is stored inside, a single block.

Example of Data Access



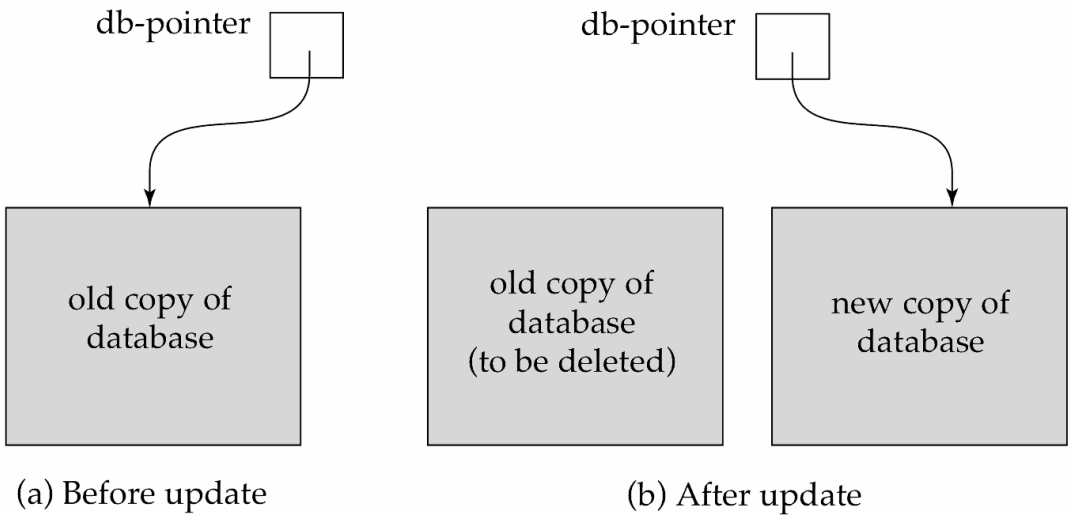
Data Access (Cont.)

- Each transaction T_i has its private work-area in which local copies of all data items accessed and updated by it are kept.
  - T_i's local copy of a data item X is called x_i.
- Transferring data items between system buffer blocks and its private work-area done by:
  - **read**(X) assigns the value of data item X to the local variable x_i.
  - **write**(X) assigns the value of local variable x_i to data item {X} in the buffer block.
  - **Note: output**(B_X) need not immediately follow **write**(X). System can perform the **output** operation when it deems fit.
- Transactions
  - Must perform **read**(X) before accessing X for the first time (subsequent reads can be from local copy)
  - **write**(X) can be executed at any time before the transaction commits

Recovery and Atomicity

- To ensure atomicity despite failures, we first output information describing the modifications to stable storage without modifying the database itself.
- We study **log-based recovery mechanisms** in detail
  - We first present key concepts
  - And then present the actual recovery algorithm
- Less used alternative: **shadow-copy** and **shadow-paging** (brief details in book)

shadow-copy



Log-Based Recovery

- A **log** is kept on stable storage.
  - The log is a sequence of **log records**, and maintains a record of update activities on the database.
- When transaction T_i starts, it registers itself by writing a
  $<T_i$ **start**>log record
- *Before* T_i executes **write**(X), a log record
  $<T_i, X, V_1, V_2>$
  is written, where V_1 is the value of X before the write (the **old value**), and V_2 is the value to be written to X (the **new value**).
- When T_i finishes it last statement, the log record $<T_i$ **commit**> is written.
- Two approaches using logs
  - Deferred database modification
  - Immediate database modification

Immediate Database Modification

- The **immediate-modification** scheme allows updates of an uncommitted transaction to be made to the buffer, or the disk itself, before the transaction commits
- Update log record must be written *before* database item is written
  - We assume that the log record is output directly to stable storage
  - (Will see later that how to postpone log record output to some extent)
- Output of updated blocks to stable storage can take place at any time before or after transaction commit
- Order in which blocks are output can be different from the order in which they are written.
- The **deferred-modification** scheme performs updates to buffer/disk only at the time of transaction commit
  - Simplifies some aspects of recovery
  - But has overhead of storing local copy

Transaction Commit

- A transaction is said to have committed when its commit log record is output to stable storage
  - all previous log records of the transaction must have been output already
- Writes performed by a transaction may still be in the buffer when the transaction commits, and may be output later

Immediate Database Modification Example

Log	Write	Output
<T_0 **start**>		
<T_0, A, 1000, 950>		
<T_0, B, 2000, 2050		
	$A = 950$	
	$B = 2050$	
<T_0 **commit**>		
<T_1 **start**>		
<T_1, C, 700, 600>		
	$C = 600$	
<T_1 **commit**>		

B_B, B_C — B_C output before T_1 commits

B_A — B_A output after T_0 commits

- Note: B_X denotes block containing X.

Concurrency Control and Recovery

- With concurrent transactions, all transactions share a single disk buffer and a single log
  - A buffer block can have data items updated by one or more transactions
- We assume that *if a transaction T_i has modified an item, no other transaction can modify the same item until T_i has committed or aborted*
  - i.e. the updates of uncommitted transactions should not be visible to other transactions
    - Otherwise how to perform undo if T1 updates A, then T2 updates A and commits, and finally T1 has to abort?
  - Can be ensured by obtaining exclusive locks on updated items and holding the locks till end of transaction (strict two-phase locking)
- Log records of different transactions may be interspersed in the log.

Undo and Redo Operations

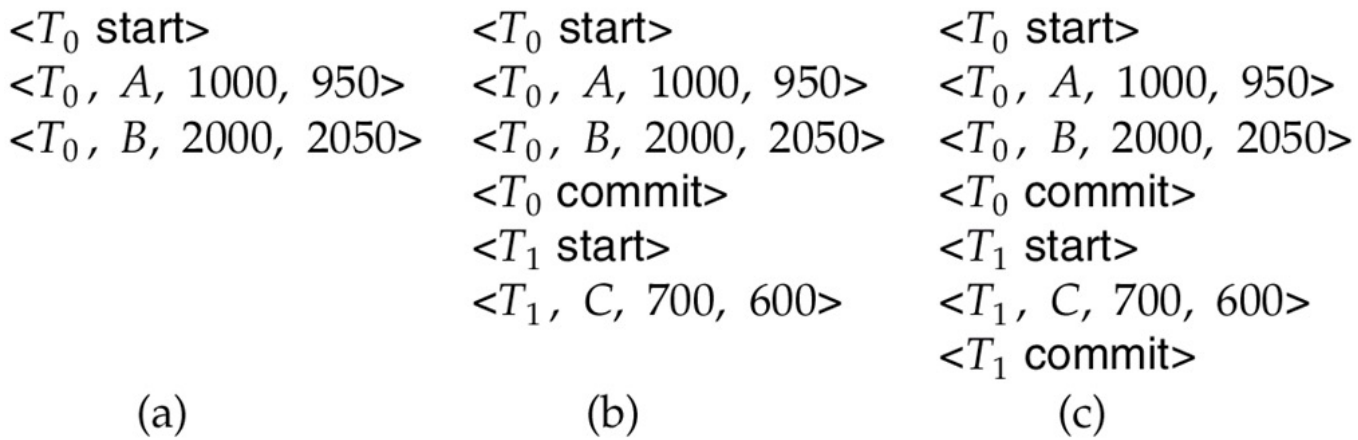
- **Undo** of a log record $<T_i, X, V_1, V_2>$ writes the **old** value V_1 to X
- **Redo** of a log record $<T_i, X, V_1, V_2>$ writes the **new** value V_2 to X
- **Undo and Redo of Transactions**
  - **undo**(T_i) restores the value of all data items updated by T_i to their old values, going backwards from the last log record for T_i
    - each time a data item X is restored to its old value V a special log record $<T_i, X, V>$ is written out
    - when undo of a transaction is complete, a log record $<T_i$ **abort**$>$ is written out.
  - **redo**(T_i) sets the value of all data items updated by T_i to the new values, going forward from the first log record for T_i
    - No logging is done in this case

Undo and Redo on Recovering from Failure

- When recovering after failure:
  - Transaction T_i needs to be undone if the log
    - contains the record $<T_i$ **start**>,
    - but does not contain either the record $<T_i$ **commit**> *or* $<T_i$ **abort**>.
  - Transaction T_i needs to be redone if the log
    - contains the records $<T_i$ **start**>
    - and contains the record $<T_i$ **commit**> *or* $<T_i$ **abort**>
- Note that If transaction T_i was undone earlier and the $<T_i$ **abort**> record written to the log, and then a failure occurs, on recovery from failure T_i is redone
  - **such a redo redoes all the original actions *including the steps that restored old values***
    - Known as **repeating history**
    - Seems wasteful, but simplifies recovery greatly

Immediate DB Modification Recovery Example

Below we show the log as it appears at three instances of time.



Recovery actions in each case above are:

(a) undo (T_0): B is restored to 2000 and A to 1000, and log records <T_0, B, 2000>, <T_0, A, 1000>, <T_0, **abort**> are written out

(b) redo (T_0) and undo (T_1): *A* and *B* are set to 950 and 2050 and C is restored to 700. Log records <T_1, C, 700>, <T_1, **abort**> are written out.

(c) redo (T_0) and redo (T_1): A and B are set to 950 and 2050 respectively. Then *C* is set to 600

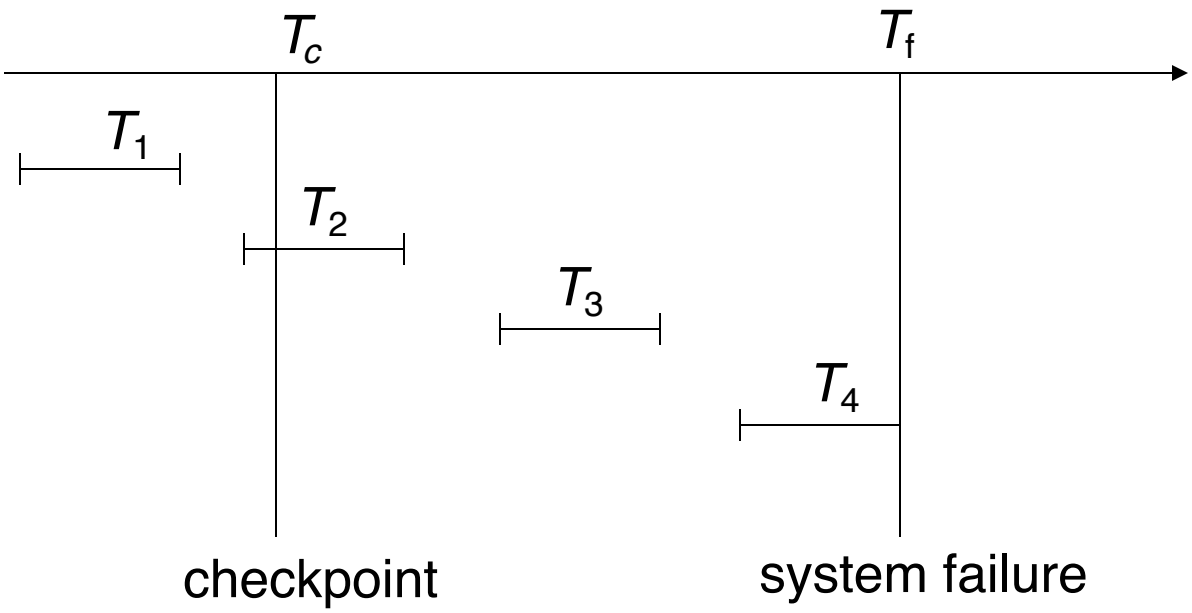
Checkpoints

- Redoing/undoing all transactions recorded in the log can be very slow
  1. processing the entire log is time-consuming if the system has run for a long time
  2. we might unnecessarily redo transactions which have already output their updates to the database.
- Streamline recovery procedure by periodically performing **checkpointing**
  1. Output all log records currently residing in main memory onto stable storage.
  2. Output all modified buffer blocks to the disk.
  3. Write a log record < **checkpoint** *L*> onto stable storage where *L* is a list of all transactions active at the time of checkpoint.
  - All updates are stopped while doing checkpointing

Checkpoints (Cont.)

- During recovery we need to consider only the most recent transaction T_i that started before the checkpoint, and transactions that started after T_i.
  1. Scan backwards from end of log to find the most recent <**checkpoint** *L*> record
     - Only transactions that are in *L* or started after the checkpoint need to be redone or undone
     - Transactions that committed or aborted before the checkpoint already have all their updates output to stable storage.
- Some earlier part of the log may be needed for undo operations
  1. Continue scanning backwards till a record <T_i **start**> is found for every transaction T_i in *L*.
     - Parts of log prior to earliest <T_i **start**> record above are not needed for recovery, and can be erased whenever desired.

Example of Checkpoints



- T_1 can be ignored (updates already output to disk due to checkpoint)
- T_2 and T_3 redone.
- T_4 undone

Recovery Algorithm

- **So far:** we covered key concepts
- **Now**: we present the components of the basic recovery algorithm
- **Later**: we present extensions to allow more concurrency

Recovery Algorithm

- **Logging** (during normal operation):
  - $<T_i$ **start**$>$ at transaction start
  - $<T_i, X_j, V_1, V_2>$ for each update, and
  - $<T_i$ **commit**$>$ at transaction end
- **Transaction rollback (during normal operation)**
  - Let T_i be the transaction to be rolled back
  - Scan log backwards from the end, and for each log record of T_i of the form $<T_i, X_j, V_1, V_2>$
    - perform the undo by writing V_1 to X_j,
    - write a log record $<T_i, X_j, V_1>$
      - such log records are called **compensation log records**
  - Once the record $<T_i$ **start**$>$ is found stop the scan and write the log record $<T_i$ **abort**$>$

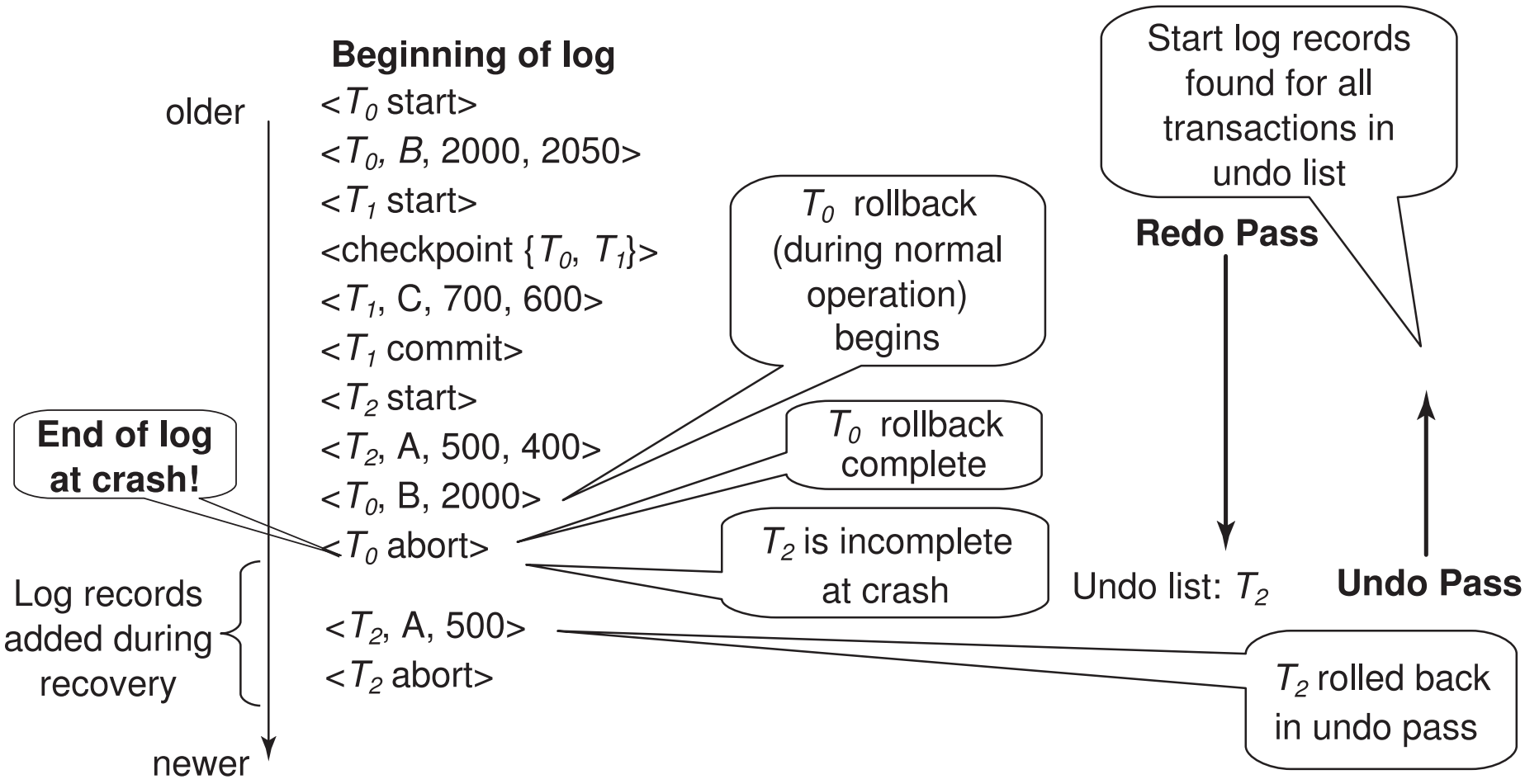
Recovery Algorithm (Cont.)

- **Recovery from failure**: Two phases
  - **Redo phase**: replay updates of **all** transactions, whether they committed, aborted, or are incomplete
  - **Undo phase**: undo all incomplete transactions
- **Redo phase**:
  1. Find last <**checkpoint** *L*> record, and set undo-list to *L*.
  2. Scan forward from above <**checkpoint** *L*> record
     1. Whenever a record $<T_i, X_j, V_1, V_2>$ or $<T_i, X_j, V_2>$ is found, redo it by writing V_2 to X_j
     2. Whenever a log record $<T_i$ **start**> is found, add T_i to undo-list
     3. Whenever a log record $<T_i$ **commit**> *or* $<T_i$ **abort**> is found, remove T_i from undo-list

Recovery Algorithm (Cont.)

- **Undo phase:**
  1. Scan log backwards from end
     1. Whenever a log record $<T_i, X_j, V_1, V_2>$ is found where T_i is in undo-list perform same actions as for transaction rollback:
        1. perform undo by writing V_1 to X_j.
        2. write a log record $<T_i, X_j, V_1>$
     2. Whenever a log record $<T_i$ **start**$>$ is found where T_i is in undo-list,
        1. Write a log record $<T_i$ **abort**$>$
        2. Remove T_i from undo-list
     3. Stop when undo-list is empty
        - i.e. $<T_i$ **start**$>$ has been found for every transaction in undo-list
- After undo phase completes, normal transaction processing can commence

Example of Recovery



Log Record Buffering

- **Log record buffering**: log records are buffered in main memory, instead of of being output directly to stable storage.
  - Log records are output to stable storage when a block of log records in the buffer is full, or a **log force** operation is executed.
- Log force is performed to commit a transaction by forcing all its log records (including the commit record) to stable storage.
- Several log records can thus be output using a single output operation, reducing the I/O cost.

Log Record Buffering (Cont.)

- The rules below must be followed if log records are buffered:
  - Log records are output to stable storage in the order in which they are created.
  - Transaction T_i enters the commit state only when the log record <T_i **commit**> has been output to stable storage.
  - Before a block of data in main memory is output to the database, all log records pertaining to data in that block must have been output to stable storage.
    - This rule is called the **write-ahead logging** or **WAL** rule
      - Strictly speaking WAL only requires undo information to be output

Database Buffering

- Database maintains an in-memory buffer of data blocks
  - When a new block is needed, if buffer is full an existing block needs to be removed from buffer
  - If the block chosen for removal has been updated, it must be output to disk
- The recovery algorithm supports the **no-force policy**: i.e., updated blocks need not be written to disk when transaction commits
  - **force policy**: requires updated blocks to be written at commit
    - More expensive commit
- The recovery algorithm supports the **steal policy**:i.e., blocks containing updates of uncommitted transactions can be written to disk, even before the transaction commits

Database Buffering (Cont.)

- If a block with uncommitted updates is output to disk, log records with undo information for the updates are output to the log on stable storage first
  - (Write ahead logging)
- No updates should be in progress on a block when it is output to disk. Can be ensured as follows.
  - Before writing a data item, transaction acquires exclusive lock on block containing the data item
  - Lock can be released once the write is completed.
    - Such locks held for short duration are called **latches**.
- **To output a block to disk**
  1. First acquire an exclusive latch on the block
     1. Ensures no update can be in progress on the block
  2. Then perform a **log flush**
  3. Then output the block to disk
  4. Finally release the latch on the block

Buffer Management (Cont.)

- Database buffer can be implemented either
  - in an area of real main-memory reserved for the database, or
  - in virtual memory
- Implementing buffer in reserved main-memory has drawbacks:
  - Memory is partitioned before-hand between database buffer and applications, limiting flexibility.
  - Needs may change, and although operating system knows best how memory should be divided up at any time, it cannot change the partitioning of memory.

Buffer Management (Cont.)

- Database buffers are generally implemented in virtual memory in spite of some drawbacks:
  - When operating system needs to evict a page that has been modified, the page is written to swap space on disk.
  - When database decides to write buffer page to disk, buffer page may be in swap space, and may have to be read from swap space on disk and output to the database on disk, resulting in extra I/O!
    - Known as **dual paging** problem.
  - Ideally when OS needs to evict a page from the buffer, it should pass control to database, which in turn should
    1. Output the page to database instead of to swap space (making sure to output log records first), if it is modified
    2. Release the page from the buffer, for the OS to use

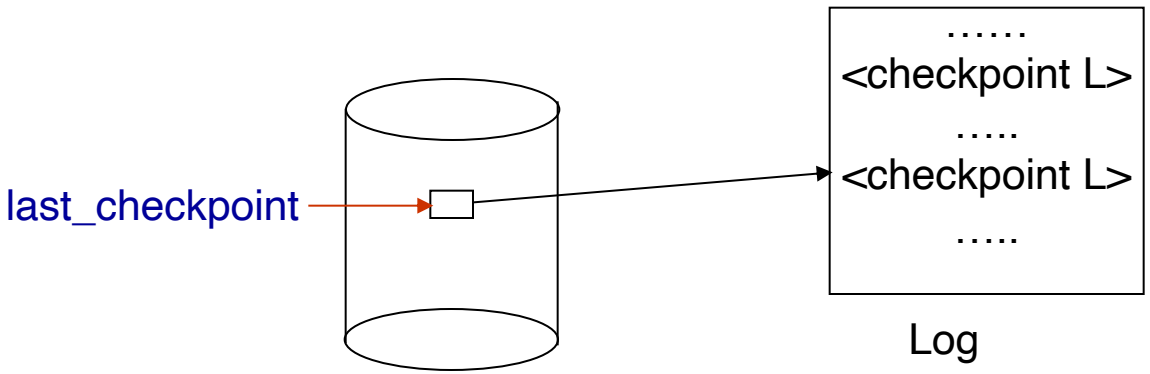
    Dual paging can thus be avoided, but common operating systems do not support such functionality.

Fuzzy Checkpointing

- To avoid long interruption of normal processing during checkpointing, allow updates to happen during checkpointing
- **Fuzzy checkpointing** is done as follows:
  1. Temporarily stop all updates by transactions
  2. Write a <**checkpoint** *L*> log record and force log to stable storage
  3. Note list *M* of modified buffer blocks
  4. Now permit transactions to proceed with their actions
  5. Output to disk all modified buffer blocks in list *M*
     - ★ blocks should not be updated while being output
     - ★ Follow WAL: all log records pertaining to a block must be output before the block is output
  6. Store a pointer to the **checkpoint** record in a fixed position **last_checkpoint** on disk

Fuzzy Checkpointing (Cont.)

- When recovering using a fuzzy checkpoint, start scan from the **checkpoint** record pointed to by **last_checkpoint**
  - Log records before **last_checkpoint** have their updates reflected in database on disk, and need not be redone.
  - Incomplete checkpoints, where system had crashed while performing checkpoint, are handled safely



Failure with Loss of Nonvolatile Storage

- So far we assumed no loss of non-volatile storage
- Technique similar to checkpointing used to deal with loss of non-volatile storage
  - Periodically **dump** the entire content of the database to stable storage
  - No transaction may be active during the dump procedure; a procedure similar to checkpointing must take place
    - Output all log records currently residing in main memory onto stable storage.
    - Output all buffer blocks onto the disk.
    - Copy the contents of the database to stable storage.
    - Output a record <**dump**> to log on stable storage.

Recovering from Failure of Non-Volatile Storage

- To recover from disk failure
  - restore database from most recent dump.
  - Consult the log and redo all transactions that committed after the dump
- Can be extended to allow transactions to be active during dump; known as **fuzzy dump** or **online dump**
  - Similar to fuzzy checkpointing

Recovery with Early Lock Release and Logical Undo Operations

Recovery with Early Lock Release

- Support for high-concurrency locking techniques, such as those used for B^+-tree concurrency control, which release locks early
  - Supports "logical undo"
- Recovery based on "repeating history", whereby recovery executes exactly the same actions as normal processing

Logical Undo Logging

- Operations like B+-tree insertions and deletions release locks early.
  - They cannot be undone by restoring old values (**physical undo**), since once a lock is released, other transactions may have updated the B+-tree.
  - Instead, insertions (resp. deletions) are undone by executing a deletion (resp. insertion) operation (known as **logical undo**).
- For such operations, undo log records should contain the undo operation to be executed
  - Such logging is called **logical undo logging**, in contrast to **physical undo logging**
    - Operations are called **logical operations**
  - Other examples:
    - delete of tuple, to undo insert of tuple
      - allows early lock release on space allocation information
    - subtract amount deposited, to undo deposit
      - allows early lock release on bank balance

Physical Redo

- Redo information is logged **physically** (that is, new value for each write) even for operations with logical undo
  - Logical redo is very complicated since database state on disk may not be “operation consistent” when recovery starts
  - Physical redo logging does not conflict with early lock release

Operation Logging

- Operation logging is done as follows:
  1. When operation starts, log <T_i, O_j, **operation-begin**>. Here O_j is a unique identifier of the operation instance.
  2. While operation is executing, normal log records with physical redo and physical undo information are logged.
  3. When operation completes, <T_i, O_j, **operation-end**, U> is logged, where U contains information needed to perform a logical undo information.

Example: insert of (key, record-id) pair (K5, RID7) into index I9

<T1, O1, operation-begin>

….

<T1, X, 10, K5>

<T1, Y, 45, RID7>

} Physical redo of steps in insert

<T1, O1, operation-end, (delete I9, K5, RID7)>

Operation Logging (Cont.)

- If crash/rollback occurs before operation completes:
  - the **operation-end** log record is not found, and
  - the physical undo information is used to undo operation.
- If crash/rollback occurs after the operation completes:
  - the **operation-end** log record is found, and in this case
  - logical undo is performed using *U*; the physical undo information for the operation is ignored.
- Redo of operation (after crash) still uses physical redo information.

Transaction Rollback with Logical Undo

Rollback of transaction T_i is done as follows:

- Scan the log backwards
  1. If a log record $<T_i, X, V_1, V_2>$ is found, perform the undo and log a al $<T_i, X, V_1>$.
  2. If a $<T_i, O_j,$ **operation-end**, $U>$ record is found
     - Rollback the operation logically using the undo information U.
       - Updates performed during roll back are logged just like during normal operation execution.
       - At the end of the operation rollback, instead of logging an **operation-end** record, generate a record

         $<T_i, O_j,$ **operation-abort**$>$.
     - Skip all preceding log records for T_i until the record $<T_i, O_j$ **operation-begin**$>$ is found

Transaction Rollback with Logical Undo (Cont.)

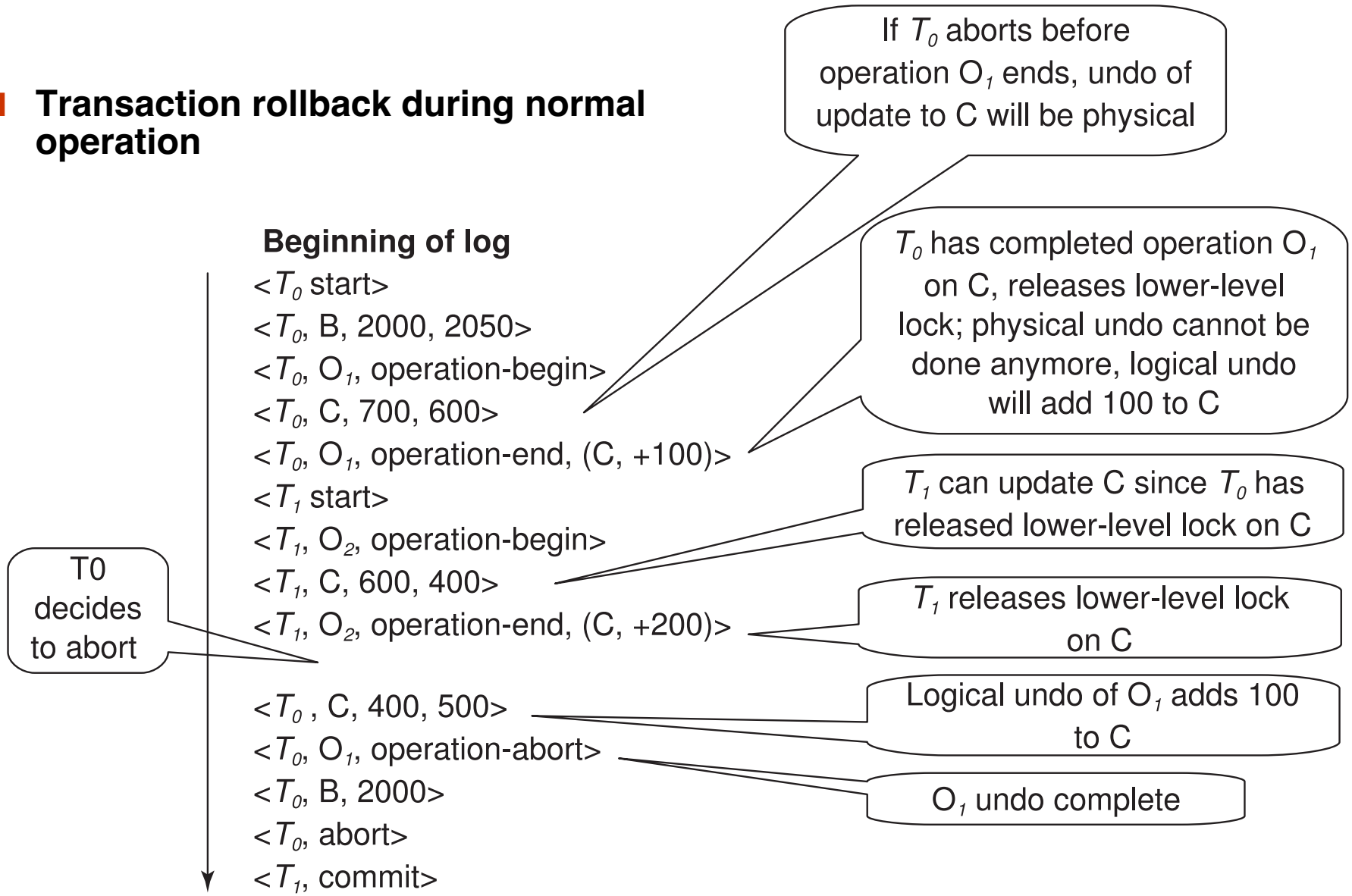
- Transaction rollback, scanning the log backwards (cont.):
  3. If a redo-only record is found ignore it
  4. If a <T_i, O_j, **operation-abort**> record is found:
     - ★ skip all preceding log records for T_i until the record <T_i, O_j, **operation-begi**n> is found.
  5. Stop the scan when the record <T_i, start> is found
  6. Add a <T_i, **abort**> record to the log

Some points to note:

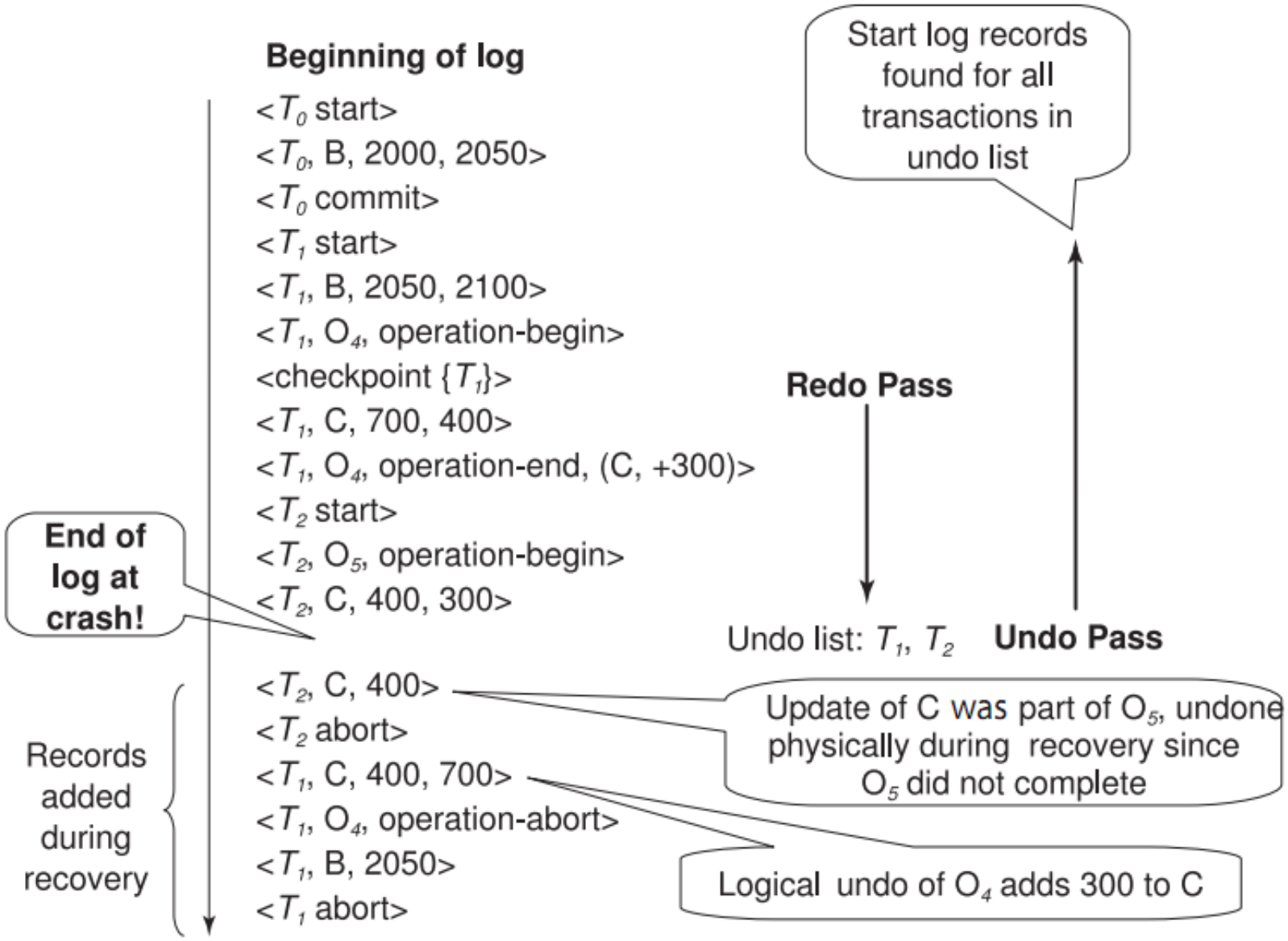
- Cases 3 and 4 above can occur only if the database crashes while a transaction is being rolled back.
- Skipping of log records as in case 4 is important to prevent multiple rollback of the same operation.

Transaction Rollback with Logical Undo

- **Transaction rollback during normal operation**



Failure Recovery with Logical Undo



Transaction Rollback: Another Example

- Example with a complete and an incomplete operation

<T1, start>

<T1, O1, operation-begin>

….

<T1, X, 10, K5>

<T1, Y, 45, RID7>

<T1, O1, operation-end, (delete I9, K5, RID7)>

<T1, O2, operation-begin>

<T1, Z, 45, 70>

← T1 Rollback begins here

<T1, Z, 45> ← redo-only log record during physical undo (of incomplete O2)

<T1, Y, .., ..> ← Normal redo records for logical undo of O1

…

<T1, O1, operation-abort> ← What if crash occurred immediately after this?

<T1, abort>

Recovery Algorithm with Logical Undo

Basically same as earlier algorithm, except for changes described earlier for transaction rollback

1. (**Redo phase**): Scan log forward from last < **checkpoint** *L*> record till end of log
   1. **Repeat history** by physically redoing all updates of all transactions,
   2. Create an undo-list during the scan as follows
      - *undo-list* is set to *L* initially
      - Whenever <T_i **start**> is found T_i is added to *undo-list*
      - Whenever <T_i **commit**> or <T_i **abort**> is found, T_i is deleted from *undo-list*

   This brings database to state as of crash, with committed as well as uncommitted transactions having been redone.

   Now *undo-list* contains transactions that are **incomplete**, that is, have neither committed nor been fully rolled back.

Recovery with Logical Undo (Cont.)

Recovery from system crash (cont.)

2. (**Undo phase**): Scan log backwards, performing undo on log records of transactions found in *undo-list*.
   - Log records of transactions being rolled back are processed as described earlier, as they are found
     - Single shared scan for all transactions being undone
   - When <T_i **start**> is found for a transaction T_i in *undo-list*, write a <T_i **abort**> log record.
   - Stop scan when <T_i **start**> records have been found for all T_i in *undo-list*

- This undoes the effects of incomplete transactions (those with neither **commit** nor **abort** log records). Recovery is now complete.

ARIES Recovery Algorithm

ARIES

- ARIES is a state of the art recovery method
  - Incorporates numerous optimizations to reduce overheads during normal processing and to speed up recovery
  - The recovery algorithm we studied earlier is modeled after ARIES, but greatly simplified by removing optimizations
- Unlike the recovery algorithm described earlier, ARIES
  1. Uses **log sequence number (LSN)** to identify log records
     - Stores LSNs in pages to identify what updates have already been applied to a database page
  2. Physiological redo
  3. Dirty page table to avoid unnecessary redos during recovery
  4. Fuzzy checkpointing that only records information about dirty pages, and does not require dirty pages to be written out at checkpoint time
     - More coming up on each of the above …

ARIES Optimizations

- **Physiological redo**
  - Affected page is physically identified, action within page can be logical
    - Used to reduce logging overheads
      - e.g. when a record is deleted and all other records have to be moved to fill hole
        - Physiological redo can log just the record deletion
        - Physical redo would require logging of old and new values for much of the page
    - Requires page to be output to disk atomically
      - Easy to achieve with hardware RAID, also supported by some disk systems
      - Incomplete page output can be detected by checksum techniques,
        - But extra actions are required for recovery
        - Treated as a media failure

ARIES Data Structures

- ARIES uses several data structures
  - Log sequence number (LSN) identifies each log record
    - Must be sequentially increasing
    - Typically an offset from beginning of log file to allow fast access
      - Easily extended to handle multiple log files
  - Page LSN
  - Log records of several different types
  - Dirty page table

ARIES Data Structures: Page LSN

- Each page contains a **PageLSN** which is the LSN of the last log record whose effects are reflected on the page
  - To update a page:
    - X-latch the page, and write the log record
    - Update the page
    - Record the LSN of the log record in PageLSN
    - Unlock page
  - To flush page to disk, must first S-latch page
    - Thus page state on disk is operation consistent
      - Required to support physiological redo
  - PageLSN is used during recovery to prevent repeated redo
    - Thus ensuring idempotence

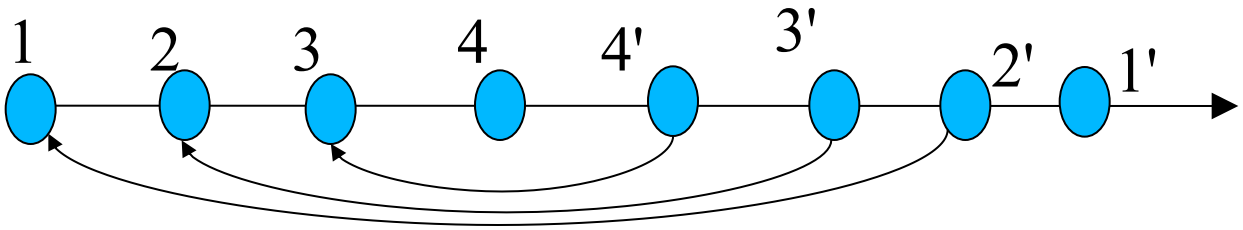
ARIES Data Structures: Log Record

- Each log record contains LSN of previous log record of the same transaction

LSN	TransID	PrevLSN	RedoInfo	UndoInfo

  - LSN in log record may be implicit
- Special redo-only log record called **compensation log record (CLR)** used to log actions taken during recovery that never need to be undone
  - Serves the role of operation-abort log records used in earlier recovery algorithm
  - Has a field UndoNextLSN to note next (earlier) record to be undone
    - Records in between would have already been undone
    - Required to avoid repeated undo of already undone actions

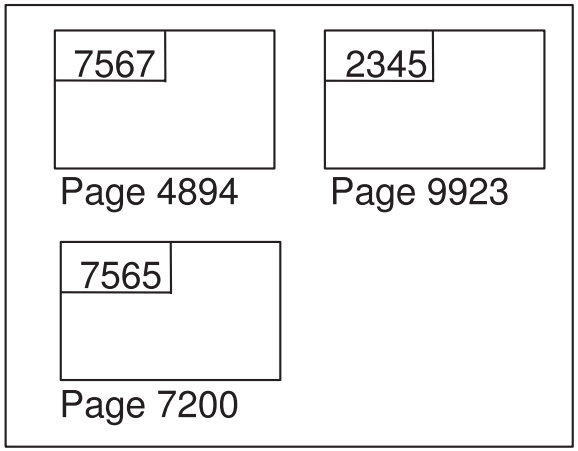
LSN	TransID	UndoNextLSN	RedoInfo



ARIES Data Structures: DirtyPage Table

- **DirtyPageTable**
  - List of pages in the buffer that have been updated
  - Contains, for each such page
    - **PageLSN** of the page
    - **RecLSN** is an LSN such that log records before this LSN have already been applied to the page version on disk
      - Set to current end of log when a page is inserted into dirty page table (just before being updated)
      - Recorded in checkpoints, helps to minimize redo work

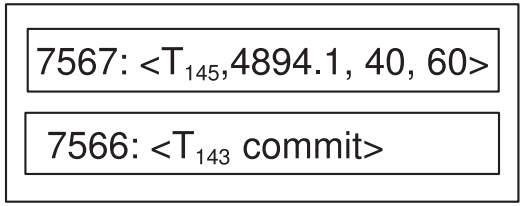
ARIES Data Structures



Database Buffer

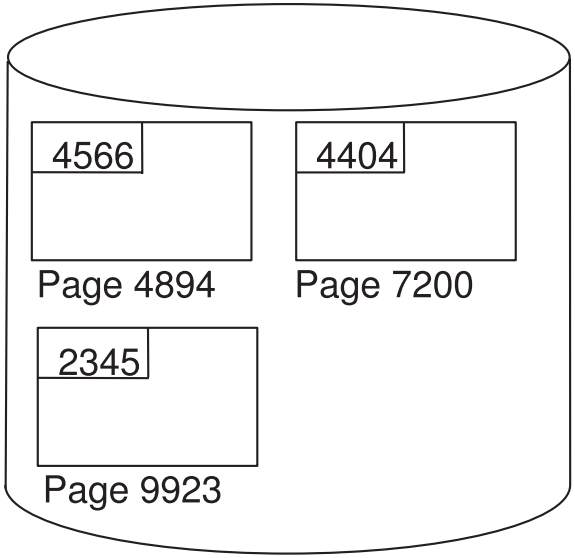
PageID	PageLSN	RecLSN
4894	7567	7564
7200	7565	7565

Dirty Page Table

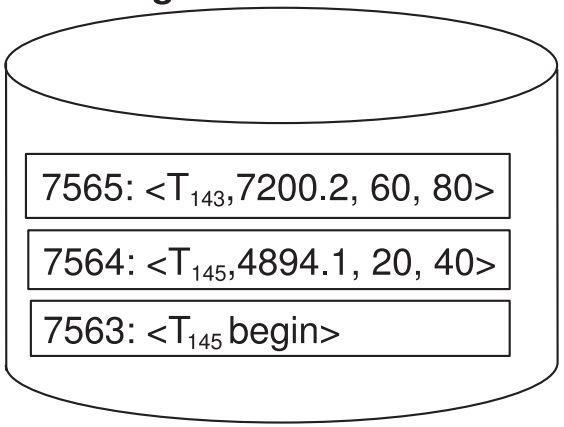


Log Buffer (PrevLSN and UndoNextLSN fields not shown)

Stable data



Stable log



ARIES Data Structures: Checkpoint Log

- **Checkpoint log record**
  - Contains:
    - DirtyPageTable and list of active transactions
    - For each active transaction, LastLSN, the LSN of the last log record written by the transaction
  - Fixed position on disk notes LSN of last completed checkpoint log record
- Dirty pages are not written out at checkpoint time
    - Instead, they are flushed out continuously, in the background
- Checkpoint is thus very low overhead
  - can be done frequently

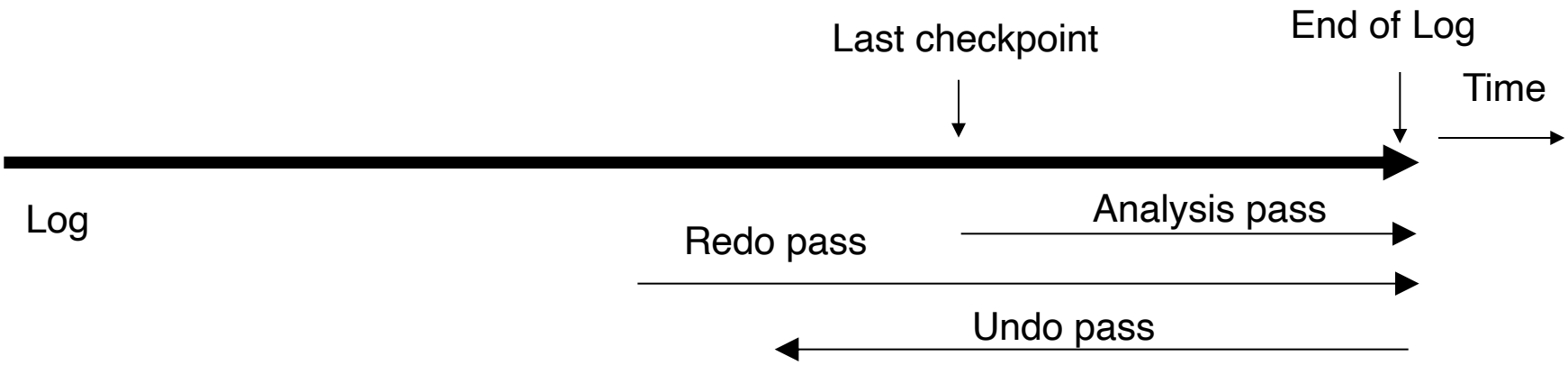
ARIES Recovery Algorithm

ARIES recovery involves three passes

- Analysis pass: Determines
  - Which transactions to undo
  - Which pages were dirty (disk version not up to date) at time of crash
  - RedoLSN: LSN from which redo should start
- Redo pass:
  - Repeats history, redoing all actions from RedoLSN
    - RecLSN and PageLSNs are used to avoid redoing actions already reflected on page
- Undo pass:
  - Rolls back all incomplete transactions
    - Transactions whose abort was complete earlier are not undone
      - Key idea: no need to undo these transactions: earlier undo actions were logged, and are redone as required

Aries Recovery: 3 Passes

- Analysis, redo and undo passes
- Analysis determines where redo should start
- Undo has to go back till start of earliest incomplete transaction



ARIES Recovery: Analysis

Analysis pass

- Starts from last complete checkpoint log record
  - Reads DirtyPageTable from log record
  - Sets RedoLSN = min of RecLSNs of all pages in DirtyPageTable
    - In case no pages are dirty, RedoLSN = checkpoint record's LSN
  - Sets undo-list = list of transactions in checkpoint log record
  - Reads LSN of last log record for each transaction in undo-list from checkpoint log record
- Scans forward from checkpoint
- .. Cont. on next page …

ARIES Recovery: Analysis (Cont.)

Analysis pass (cont.)

- Scans forward from checkpoint
  - If any log record found for transaction not in undo-list, adds transaction to undo-list
  - Whenever an update log record is found
    - If page is not in DirtyPageTable, it is added with RecLSN set to LSN of the update log record
  - If transaction end log record found, delete transaction from undo-list
  - Keeps track of last log record for each transaction in undo-list
    - May be needed for later undo
- At end of analysis pass:
  - RedoLSN determines where to start redo pass
  - RecLSN for each page in DirtyPageTable used to minimize redo work
  - All transactions in undo-list need to be rolled back

ARIES Redo Pass

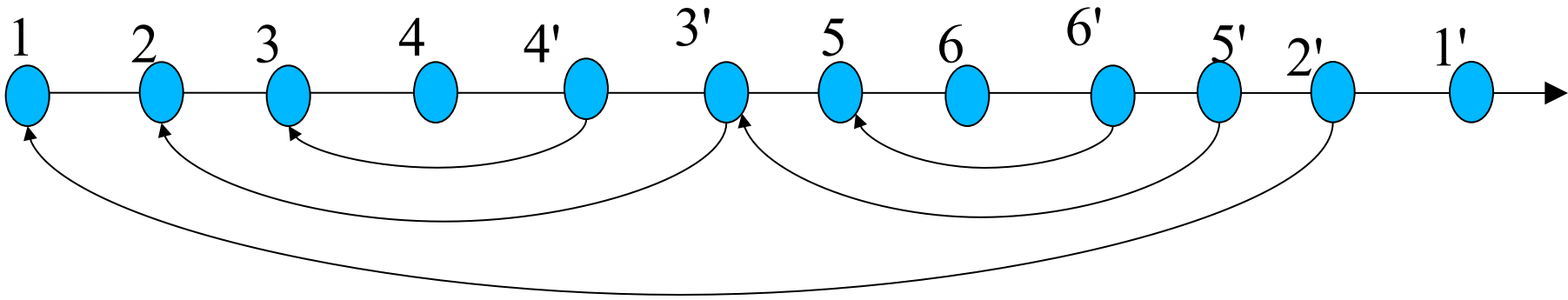
Redo Pass: Repeats history by replaying every action not already reflected in the page on disk, as follows:

- Scans forward from RedoLSN. Whenever an update log record is found:
  1. If the page is not in DirtyPageTable or the LSN of the log record is less than the RecLSN of the page in DirtyPageTable, then skip the log record
  2. Otherwise fetch the page from disk. If the PageLSN of the page fetched from disk is less than the LSN of the log record, redo the log record

  NOTE: if either test is negative the effects of the log record have already appeared on the page. First test avoids even fetching the page from disk!

ARIES Undo Actions

- When an undo is performed for an update log record
  - Generate a CLR containing the undo action performed (actions performed during undo are logged physicaly or physiologically).
    - CLR for record *n* noted as *n*' in figure below
  - Set UndoNextLSN of the CLR to the PrevLSN value of the update log record
    - Arrows indicate UndoNextLSN value
- ARIES supports partial rollback
  - Used e.g. to handle deadlocks by rolling back just enough to release reqd. locks
  - Figure indicates forward actions after partial rollbacks
    - records 3 and 4 initially, later 5 and 6, then full rollback

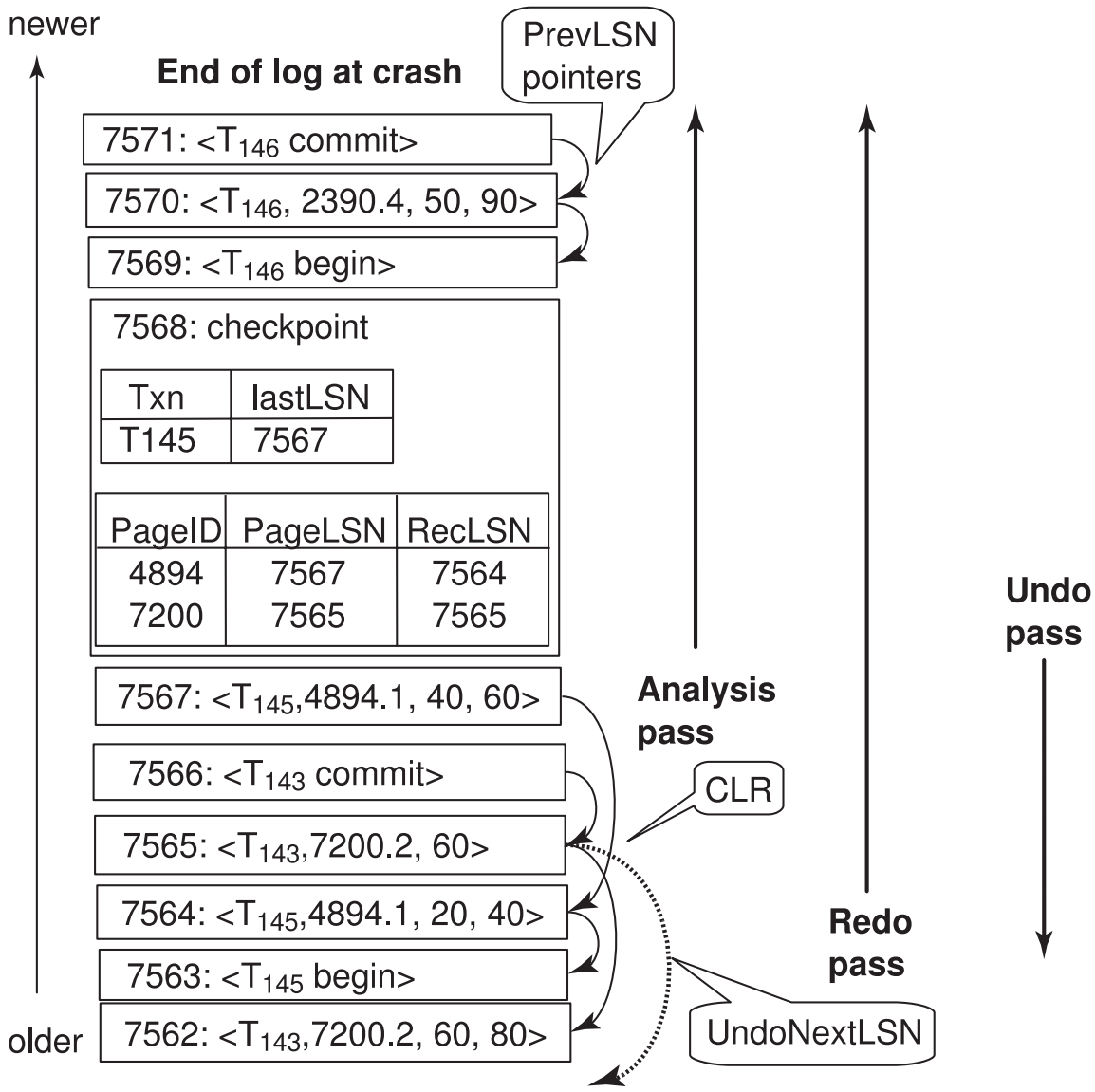


ARIES: Undo Pass

Undo pass:

- Performs backward scan on log undoing all transaction in undo-list
  - Backward scan optimized by skipping unneeded log records as follows:
    - Next LSN to be undone for each transaction set to LSN of last log record for transaction found by analysis pass.
    - At each step pick largest of these LSNs to undo, skip back to it and undo it
    - After undoing a log record
      - For ordinary log records, set next LSN to be undone for transaction to PrevLSN noted in the log record
      - For compensation log records (CLRs) set next LSN to be undo to UndoNextLSN noted in the log record
        - All intervening records are skipped since they would have been undone already
- Undos performed as described earlier

Recovery Actions in ARIES



Other ARIES Features

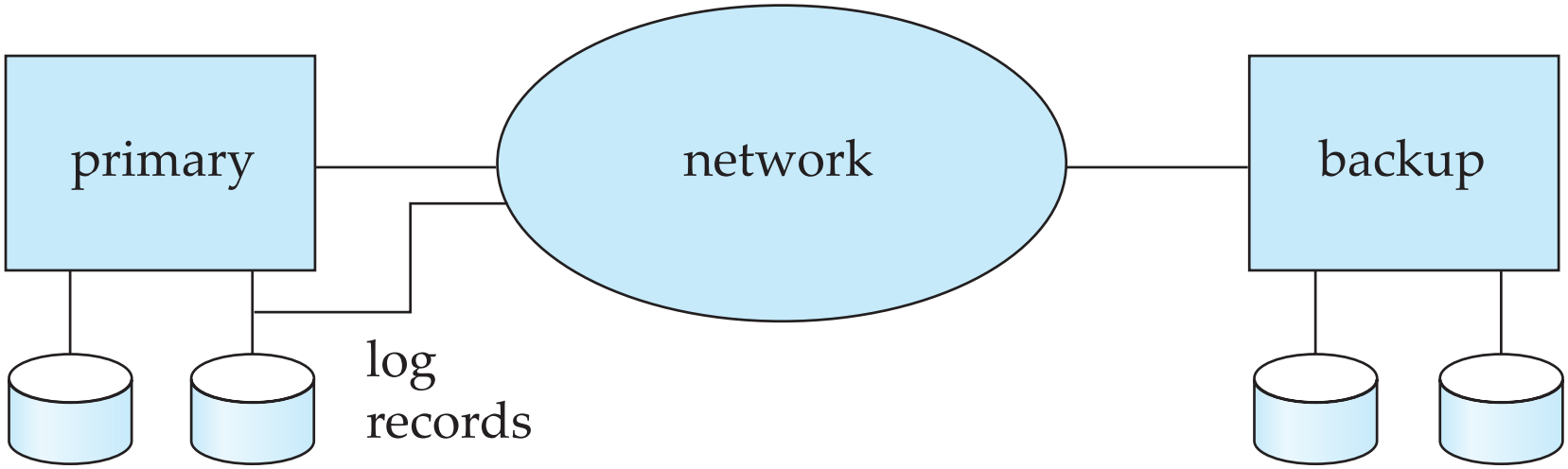
- Recovery Independence
  - Pages can be recovered independently of others
    - E.g. if some disk pages fail they can be recovered from a backup while other pages are being used
- Savepoints:
  - Transactions can record savepoints and roll back to a savepoint
    - Useful for complex transactions
    - Also used to rollback just enough to release locks on deadlock

Other ARIES Features (Cont.)

- Fine-grained locking:
  - Index concurrency algorithms that permit tuple level locking on indices can be used
    - These require logical undo, rather than physical undo, as in earlier recovery algorithm
- Recovery optimizations: For example:
  - Dirty page table can be used to prefetch pages during redo
  - Out of order redo is possible:
    - redo can be postponed on a page being fetched from disk, and
      performed when page is fetched.
    - Meanwhile other log records can continue to be processed

Remote Backup Systems

- Remote backup systems provide high availability by allowing transaction processing to continue even if the primary site is destroyed.



Remote Backup Systems (Cont.)

- **Detection of failure**: Backup site must detect when primary site has failed
  - to distinguish primary site failure from link failure maintain several communication links between the primary and the remote backup.
  - Heart-beat messages
- **Transfer of control**:
  - To take over control backup site first perform recovery using its copy of the database and all the long records it has received from the primary.
    - Thus, completed transactions are redone and incomplete transactions are rolled back.
  - When the backup site takes over processing it becomes the new primary
  - To transfer control back to old primary when it recovers, old primary must receive redo logs from the old backup and apply all updates locally.

Remote Backup Systems (Cont.)

- **Time to recover**: To reduce delay in takeover, backup site periodically proceses the redo log records (in effect, performing recovery from previous database state), performs a checkpoint, and can then delete earlier parts of the log.
- **Hot-Spare** configuration permits very fast takeover:
  - Backup continually processes redo log record as they arrive, applying the updates locally.
  - When failure of the primary is detected the backup rolls back incomplete transactions, and is ready to process new transactions.
- Alternative to remote backup: distributed database with replicated data
  - Remote backup is faster and cheaper, but less tolerant to failure
    - more on this in Chapter 19

Remote Backup Systems (Cont.)

- Ensure durability of updates by delaying transaction commit until update is logged at backup; avoid this delay by permitting lower degrees of durability.
- **One-safe:** commit as soon as transaction's commit log record is written at primary
  - Problem: updates may not arrive at backup before it takes over.
- **Two-very-safe:** commit when transaction's commit log record is written at primary and backup
  - Reduces availability since transactions cannot commit if either site fails.
- **Two-safe:** proceed as in two-very-safe if both primary and backup are active. If only the primary is active, the transaction commits as soon as is commit log record is written at the primary.
  - Better availability than two-very-safe; avoids problem of lost transactions in one-safe.