Implementing Isolation

Chapter 20

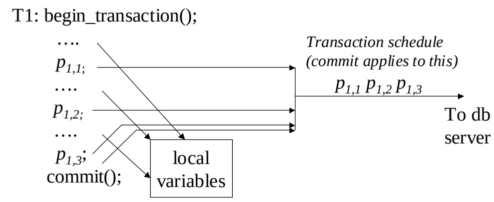
The Issue

- Maintaining database correctness when many transactions are accessing the database concurrently
  - – Assuming each transaction maintains database correctness when executed in isolation

Isolation

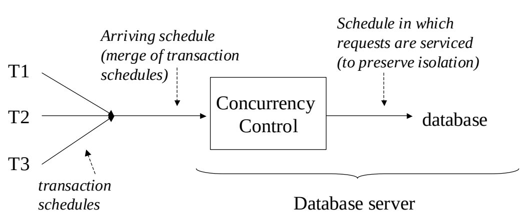
- Serial execution:
  - – Since each transaction is consistent and isolated from all others, schedule is guaranteed to be correct for all applications
  - – Inadequate performance
    - Since system has multiple asynchronous resources and transaction uses only one at a time
- Concurrent execution:
  - – Improved performance (multiprogramming)
  - – Some interleavings produce correct result, others do not
  - – We are interested in concurrent schedules that are *equivalent* to serial schedules. These are referred to as *serializable* schedules.

Transaction Schedule



- Consistent - performs correctly when executed in isolation starting in a consistent database state
  - – Preserves database consistency
  - – Moves database to a new state that corresponds to new real-world state

Schedule



Schedule

- Representation 1:

  T_1: p_1 p_2 p_3 p_4
  T_2: p_1 p_2
  time $\rightarrow$

- Representation 2:

  $p_{1,1}$ $p_{1,2}$ $p_{2,1}$ $p_{1,3}$ $p_{2,2}$ $p_{1,4}$
  time $\rightarrow$

Concurrency Control

- Transforms arriving interleaved schedule into a correct interleaved schedule to be submitted to the DBMS
  - Delays servicing a request (reordering) - causes a transaction to wait
  - Refuses to service a request - causes transaction to abort
- Actions taken by concurrency control have performance costs
  - Goal is to avoid delaying or refusing to service a request

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Correct Schedules

- Interleaved schedules *equivalent* to serial schedules are the only ones guaranteed to be correct for *all* applications
- Equivalence based on *commutativity* of operations
- **Definition:** Database operations p_1 and p_2 commute if, *for all initial database states*, they
  (1) *return the same results* and
  (2) *leave the database in the same final state*
  when executed in either order.

$p_1 \ p_2 \qquad p_2 \ p_1$

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Conventional Operations

- Read
  - *r(x, X)* - copy the value of database variable *x* to local variable *X*
- Write
  - *w(x, X)* - copy the value of local variable *X* to database variable *x*
- We use $r_1(x)$ and $w_1(x)$ to mean a read or write of *x* by transaction T_1

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Commutativity of Read and Write Operations

- p_1 commutes with p_2 if
  - They operate on different data items
    - $w_1(x)$ commutes with $w_2(y)$ and $r_2(y)$
  - Both are reads
    - $r_1(x)$ commutes with $r_2(x)$
- Operations that do not commute *conflict*
    - $w_1(x)$ conflicts with $w_2(x)$
    - $w_1(x)$ conflicts with $r_2(x)$

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Equivalence of Schedules

- An interchange of adjacent operations *of different transactions* in a schedule creates an equivalent schedule if the operations commute

  S_1: $S_{1,1}\ p_{i,j}\ p_{k,l}\ S_{1,2}$ where $i \neq k$
  S_2: $S_{1,1}\ p_{k,l}\ p_{i,j}\ S_{1,2}$
  - Each transaction computes the same results (since operations return the same values in both schedules) and hence writes the same values to the database.
  - The database is left in the same final state (since the state seen by $S_{1,2}$ is the same in both schedules).

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Equivalence of Schedules

- Equivalence is transitive: If S_1 can be derived from S_2 by a series of such interchanges, S_1 is equivalent to S_2

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Example of Equivalence

conflict

S_1: $r_1(x)\ r_2(x)\ w_2(x)\ r_1(y)\ w_1(y)$

S_2: $r_1(x)\ r_2(x)\ r_1(y)\ w_2(x)\ w_1(y)$

S_3: $r_1(x)\ r_1(y)\ r_2(x)\ w_2(x)\ w_1(y)$

S_4: $r_1(x)\ r_1(y)\ r_2(x)\ w_1(y)\ w_2(x)$

S_5: $r_1(x)\ r_1(y)\ w_1(y)\ r_2(x)\ w_2(x)$

conflicting operations ordered in same way

S_1 is equivalent to S_5
S_5 is the serial schedule T_1, T_2
S_1 is *serializable*
S_1 is *not* equivalent to the serial schedule T_2, T_1

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Example of Equivalence

T_1: *begin transaction*
read (x, X);
X = X+4;
write (x, X);
commit;

T_2: *begin transaction*
read (x,Y);
write (y,Y);
commit;

	<u>*initial state*</u>		<u>*final state*</u>
	$x=1, y=3$	$r_1(x)\ r_2(x)\ w_2(y)\ w_1(x)$	$x=5, y=1$
Interchange commuting operations	$x=1, y=3$	$r_2(x)\ w_2(y)\ r_1(x)\ w_1(x)$ T_2 T_1	$x=5, y=1$
Interchange conflicting operations	$x=1, y=3$	$r_1(x)\ w_1(x)\ r_2(x)\ w_2(y)$ T_1 T_2	$x=5, y=5$

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Serializable Schedules

- S is serializable if it is equivalent to a serial schedule
- Transactions are totally isolated in a serializable schedule
- A schedule is correct for *any* application if it is a serializable schedule of consistent transactions
- The schedule :
  $r_1(x)\ r_2(y)\ w_2(x)\ w_1(y)$
  is *not* serializable

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Isolation Levels

- Serializability provides a *conservative* definition of correctness
  - For a particular application there might be many acceptable *non*-serializable schedules
  - Requiring serializability might degrade performance
- DBMSs offer a variety of isolation levels:
  - SERIALIZABLE is the most stringent
  - Lower levels of isolation give better performance
    - *Might* allow incorrect schedules
    - *Might* be adequate for some applications

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Serializable

- **Theorem** - Schedule S_1 can be derived from S_2 by a sequence of commutative interchanges if and only if conflicting operations in S_1 and S_2 are ordered in the same way

  Only If: Commutative interchanges do not reorder conflicting operations

  If: A sequence of commutative interchanges can be determined that takes S_1 to S_2 since conflicting operations do not have to be reordered (see text)

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Conflict Equivalence

- **Definition**- Two schedules, S_1 and S_2, of the same set of operations are *conflict equivalent* if conflicting operations are ordered in the same way in both
  - Or (using theorem) if one can be obtained from the other by a series of commutative interchanges

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Conflict Equivalence

- **Result**- A schedule is serializable if it is conflict equivalent to a serial schedule

$$\underbrace{r_1(x)\ w_2(x)}_{conflict}\ \underbrace{w_1(y)\ r_2(y)}_{conflict} \equiv r_1(x)\ w_1(y)\ w_2(x)\ r_2(y)$$

- If in S transactions T_1 and T_2 have several pairs of conflicting operations ($p_{1,1}$ conflicts with $p_{2,1}$ and $p_{1,2}$ conflicts with $p_{2,2}$) then $p_{1,1}$ must precede $p_{2,1}$ and $p_{1,2}$ must precede $p_{2,2}$ (or vice versa) in order for S to be serializable.

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View Equivalence

- Two schedules of the same set of operations are *view equivalent* if:
  - Corresponding read operations in each return the same values (hence computations are the same)
  - Both schedules yield the same final database state
- Conflict equivalence implies view equivalence.
- View equivalence *does not* imply conflict equivalence.

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View Equivalence

T_1: $w(y)$ $w(x)$
T_2: $r(y)$ $w(x)$
T_3: $w(x)$

T_1:		$w(y)$	$w(x)$		
T_2:	$r(y)$			$w(x)$	
T_3:					$w(x)$

- Schedule *is not* conflict equivalent to a serial schedule
- Schedule has same effect as serial schedule $T_2\ T_1\ T_3$. It *is* view equivalent to a serial schedule and hence serializable

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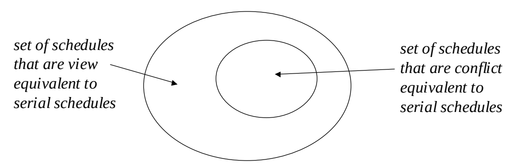
Conflict vs View Equivalence



- A concurrency control based on view equivalence should provide better performance than one based on conflict equivalence since less reordering is done but …
- It is difficult to implement a view equivalence concurrency control

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Conflict Equivalence and Serializability

- Serializability is a conservative notion of correctness and conflict equivalence provides a conservative technique for determining serializability
- However, a concurrency control that guarantees conflict equivalence to serial schedules ensures correctness and is easily implemented

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Serialization Graph of a Schedule, S

- Nodes represent transactions
- There is a directed edge from node T_i to node T_j if T_i has an operation $p_{i,k}$ that conflicts with an operation $p_{j,r}$ of T_j and $p_{i,k}$ precedes $p_{j,r}$ in S
- **Theorem** - A schedule is conflict serializable if and only if its serialization graph has no cycles

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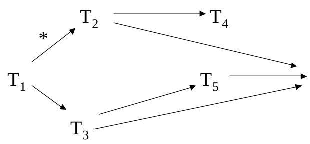
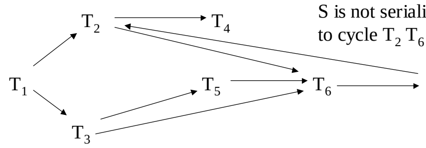
Example

Conflict (*)

S: … $p_{1,i}$, …, $p_{2,j}$, …



S is serializable in order $T_1\ T_2\ T_3\ T_4\ T_5\ T_6\ T_7$



S is not serializable due to cycle $T_2\ T_6\ T_7\ T_2$

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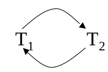
Intuition: Serializability and Nonserializability

- Consider the nonserializable schedule

  $r_1(x)\ w_2(x)\ r_2(y)\ w_1(y)$



- Two ways to think about it:
  - Because of the conflicts, the operations of T_1 and T_2 cannot be interchanged to make an equivalent serial schedule
  - Because T_1 read *x* before T_2 wrote it, T_1 must precede T_2 in any ordering, and because T_1 wrote *y* after T_2 read it, T_1 must follow T_2 in any ordering --- clearly an impossibility

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Recoverability: Schedules with Aborted Transactions

T_1 : *r (x) w(y) commit*
T_2: *w(x)* *abort*

- T_2 has aborted but has had an indirect effect on the database – schedule is *un*recoverable
- **Problem**: T_1 read uncommitted data - *dirty read*
- **Solution:** A concurrency control is *recoverable* if it does not allow T_1 to commit until all other transactions that wrote values T_1 read have committed

T_1 : *r (x) w(y)* request commit *abort*
T_2: *w(x)* *abort*

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Cascaded Abort

- Recoverable schedules solve abort problem but allow *cascaded abort*: abort of one transaction forces abort of another

T_1: *r (y) w(z)* *abort*
T_2: *r (x) w(y)* *abort*
T_3: *w(x)* *abort*

- Better solution: prohibit dirty reads

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Dirty Write

- *Dirty write*: A transaction writes a data item written by an active transaction
- Dirty write complicates rollback:

no rollback necessary

T_1: *w(x)* *abort*
T_2 : *w(x)* *abort*

what value of x should be restored?

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Strict Schedules

- *Strict schedule*: Dirty writes and dirty reads are prohibited
- Strict and serializable are two different properties
  - Strict, non-serializable schedule:

    $r_1(x)\ w_2(x)\ r_2(y)\ w_1(y)\ c_1\ c_2$
  - Serializable, non-strict schedule:

    $w_2(x)\ r_1(x)\ w_2(y)\ r_1(y)\ c_1\ c_2$

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Concurrency Control

- Concurrency control cannot see entire schedule:
  - It sees one request at a time and must decide whether to allow it to be serviced
- Strategy: Do not service a request if:
  - It violates strictness or serializability, or
  - There is a possibility that a subsequent arrival might cause a violation of serializability

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Models of Concurrency Controls

- **Immediate Update** – (*the model we have discussed*)
  - A write updates a database item
  - A read copies value from a database item
  - Commit makes updates durable
  - Abort undoes updates
- **Deferred Update –** *(we will discuss this later)*
  - A write stores new value in the transaction's intentions list (does *not* update the database)
  - A read copies value from the database or the transaction's intentions list
  - Commit uses intentions list to durably update database
  - Abort discards intentions list

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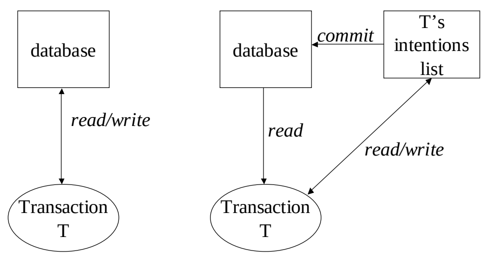
Immediate vs. Deferred Update



Immediate Update

Deferred Update

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Models of Concurrency Controls

- **Pessimistic** –
  - A transaction requests permission for each database (read/write) operation
  - Concurrency control can:
    - *Grant* the operation (submit it for execution)
    - *Delay* it until a subsequent event occurs (commit or abort of another transaction), or
    - *Abort* the transaction
  - Decisions are made *conservatively* so that a commit request can *always* be granted
    - Takes precautions even if conflicts do not occur

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Models of Concurrency Controls

- **Optimistic -**
  - Request for database operations (read/write) are *always granted*
  - Request to commit *might be denied*
    - Transaction is aborted if it performed a non-serializable operation
  - Assumes that conflicts are not likely

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Immediate-Update Pessimistic Control

- The most commonly used control
- Consider first a simple case
  - Suppose such a control allowed a transaction, T_1 , to perform some operation and then, while T_1 was still active ,it allowed another transaction, T_2 , to perform a conflicting operation
  - The schedule would not be strict and so this situation cannot be allowed
    - But consider a bit further what might happen …

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Immediate-Update Pessimistic Control

- If T_1 executes $op_1(x)$ and then T_2 executes a conflicting operation, $op_2(x)$, T_2 must follow T_1 in any equivalent serial schedule.
- **Problem**: If T_1 and T_2 later make conflicting accesses to *y*, control cannot allow ordering $op'_2(y), op'_1(y)$
  - control has to use transitive closure of transaction ordering to prevent loop in serialization graph (too complicated)
- **Worse problem**:

$$\underbrace{w_1(x)\ r_2(x)\ w_2(y)\ commit_2}_{\text{looks good}}\ \ \underset{\substack{\uparrow\\ \text{disaster}}}{request\_r_1(y)}$$

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Immediate-Update Pessimistic Control

- **Rule:**
  - *Do not grant* a request that imposes an ordering among *active* transactions (*delay* the requesting transaction)
  - *Grant* a request that does not conflict with previously granted requests of *active* transactions
- Rule can be used as each request arrives
- If a transaction's request is delayed, it is forced to wait (but the transaction is still considered active)
  - Delayed requests are reconsidered when a transaction completes (aborts or commits) since it becomes inactive

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Immediate-Update Pessimistic Control

- **Result**: Each schedule, S, is equivalent to a serial schedule in which transactions are ordered in the order in which they commit in S (and possibly other serial schedules as well)
  - **Reason**: When a transaction commits, none of its operations conflict with those of other active transactions. Therefore it can be ordered before all active transactions.
  - **Example**: The following (non-serializable) schedule is not permitted because T_1 was active at the time $w_2(x)$ (which conflicts with $r_1(x)$) was requested

$$r_1(x)\ \ w_2(x)\ \ r_2(y)\ \ w_1(y)$$

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Immediate-Update Pessimistic Control

$$S:\ \underbrace{op_1\ op_2 \ldots op_n}_{\text{no conflicting operations}}\ \underset{\substack{\uparrow\\ \text{first commit}}}{c_1}$$

$$S':\ \underset{\substack{\uparrow\\ \text{all operations of } T_1}}{T_1}\ \ \underbrace{op'_1\ op'_2 \ldots op'_n}_{\text{remaining operations of } S}$$

- S and S′ are conflict equivalent
  - The argument can be repeated at subsequent commits

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Immediate-Update Pessimistic Control

- Commit order is useful since transactions might perform external actions visible to users
  - After a deposit transaction commits, you expect a subsequent transaction to see the new account balance

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Deadlock

- **Problem**: Controls that cause transactions to wait can cause deadlocks

$$w_1(x)\ w_2(y)\ \ \substack{\text{request}\\ r_1(y)}\ \ \substack{\text{request}\\ r_2(x)}$$

- **Solution**: Abort one transaction in the cycle
  - Use wait-for graph to detect cycle when a request is delayed or
  - Assume a deadlock when a transaction waits longer than some time-out period

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Locking Implementation of an Immediate-Update Pessimistic Control

- A transaction can read a database item if it holds a read (shared) lock on the item
- It can read *or* update the item if it holds a write (exclusive) lock
- If the transaction does not already hold the required lock, a lock request is automatically made as part of the (read or write) request

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Locking

- Request for read lock on an item is granted if no transaction currently holds write lock on the item
  - Cannot read an item written by an active transaction
- Request for write lock granted if no transaction holds any lock on item
  - Cannot write an item read/written by an active transaction
- Transaction is delayed if request cannot be granted

Requested mode	Granted mode *read*	Granted mode *write*
read		x
write	x	x

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Locking

- All locks held by a transaction are released when the transaction completes (commits or aborts)
  - Delayed requests are re-examined at this time

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Locking

- **Result**: A lock is not granted if the requested access conflicts with a prior access of an active transaction; instead the transaction waits. This enforces the rule:
  - Do not grant a request that imposes an ordering among active transactions (delay the requesting transaction)
- Resulting schedules are serializable and strict

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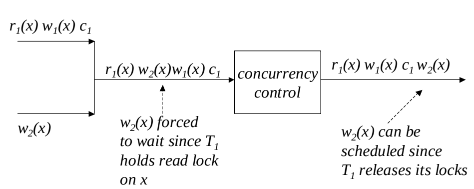
Locking



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Locking Implementation

- Associate a *lock set, L(x),* and a *wait set, W(x),* with each active database item, *x*
  - *L(x)* contains an entry for each granted lock on *x*
  - *W(x)* contains an entry for each pending request on *x*
  - When an entry is removed from *L(x)* (due to transaction termination), promote (non-conflicting) entries from *W(x)* using some scheduling policy (*e.g.*, FCFS)
- Associate a lock list, $\mathcal{L}_i$, with each transaction, T_i.
  - $\mathcal{L}_i$ links T_i's elements in all lock and wait sets
  - Used to release locks on termination

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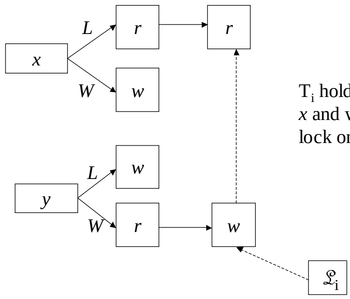
Locking Implementation



T_i holds an *r* lock on *x* and waits for a *w* lock on *y*

Manual Locking

- Better performance possible if transactions are allowed to release locks before commit
  - Ex: release lock on item when finished accessing the item

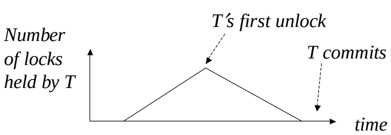
T_1: *l(x) r(x) l(y) r(y) u(x) w(y) u(y)*
T_2: *l(x) l(z) w(x) w(z) u(x) u(z)*

- However, early lock release can lead to non-serializable schedules

T_1: *l(x) r(x) u(x)* *l(y) r(y) u(y)*
T_2: *l(x) l(y) w(x) w(y)* $\underbrace{u(x)\ u(y)}_{commit}$

Two-Phase Locking

- Transaction does not release a lock until it has all the locks it will ever require.
- Transaction has a locking phase followed by an unlocking phase



- Guarantees serializability when locking is done manually

Two-Phase Locking

- **Theorem**: A concurrency control that uses two phase locking produces only serializable schedules.
  - *Proof (sketch):* Consider two transactions T_1 and T_2 in schedule S produced by a two-phase locking control and assume T_1's first unlock, t_1, precedes T_2's first unlock, t_2.
    - If they do not access common data items, then all operations commute.
    - Suppose they do. All of T_1's accesses to common items precede all of T_2's. If this were not so, T_2's first unlock must precede a lock request of T_1. Since both transactions are two-phase, this implies that T_2's first unlock precedes T_1's first unlock, contradicting the assumption. Hence, all conflicts between T_1 and T_2 are in the same direction.
  - It follows that the serialization graph is cycle-free since if there is a cycle T_1, T_2, …T_n then it must be the case that $t_1 < t_2 < \ldots < t_n < t_1$

Two-Phase Locking

- A schedule produced by a two-phase locking control is:
  - Equivalent to a serial schedule in which transactions are ordered by the time of their first unlock operation
  - Not necessarily recoverable (dirty reads and writes are possible)

T1: *l(x) r(x) l(y) w(y) u(y)* *abort*
T2: *l(y) r(y) l(z) w(z) u(z) u(y) commit*

Two-Phase Locking

- A two-phase locking control that holds write locks until commit produces strict, serializable schedules
- A strict two-phase locking control holds *all* locks until commit and produces strict serializable schedules
  - This is automatic locking
  - Equivalent to a serial schedule in which transactions are ordered by their commit time
- "Strict" is used in two different ways: a control that releases read locks early guarantees *strictness*, but is not *strict* two-phase locking control

Lock Granularity

- Data item: variable, record, row, table, file
- When an item is accessed, the DBMS locks an entity that *contains* the item. The size of that entity determines the *granularity* of the lock
  - Coarse granularity (large entities locked)
    - **Advantage**: If transactions tend to access multiple items in the same entity, fewer lock requests need to be processed and less lock storage space required
    - **Disadvantage:** Concurrency is reduced since some items are unnecessarily locked
  - Fine granularity (small entities locked)
    - Advantages and disadvantages are reversed

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Lock Granularity

- Table locking (*coarse*)
  - Lock entire table when a row is accessed.
- Row (tuple) locking (*fine*)
  - Lock only the row that is accessed.
- Page locking (compromise)
  - When a row is accessed, lock the containing page

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Objects and Semantic Commutativity

- Read/write operations have little associated semantics and hence little associated commutativity.
  - Among operations on the same item, only reads commute.
- Abstract operations (for example operations on objects) have more semantics, allowing
  - More commutativity to be recognized
  - More concurrency to be achieved

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Abstract Operations and Commutativity

- A concurrency control that deals with operations at an abstract level can recognize more commutativity and achieve more concurrency
- **Example**: operations *deposit(acct,n)*, *withdraw(acct,n)* on an account object (where *n* is the dollar amount)

	Granted Mode	
Requested Mode	*deposit( )*	*withdraw( )*
deposit( )		X
withdraw( )	X	X

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A Concurrency Control Based on Abstract Operations

- Concurrency control grants *deposit* and *withdraw* locks based on this table
- If one transaction has a *deposit* lock on an account object, another transaction can also obtain a *deposit* lock on the object
- Would not be possible if control viewed *deposit* as a *read* followed by a *write* and attempted to get *read* and *write* locks

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A Concurrency Control Based on Abstract Operations

- Since T_1 and T_2 can both hold a *deposit* lock on the same *account* object their deposit operations do not delay each other
  - As a result, the schedule can contain:
    … $deposit_1(acct,n)$ … $deposit_2(acct,m)$ … $commit_1$
    or
    … $deposit_2(acct,m)$ … $deposit_1(acct,n)$ … $commit_2$
  - But the two deposit operations must be isolated from each other. Assuming *bal* is the account balance, the schedule
    $r_1(bal)\ r_2(bal)\ w_1(bal)\ w_2(bal)$
    cannot be allowed

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Partial vs. Total Operations

- *deposit( ), withdraw( )* are *total operations*: they are defined in all database states.
- *withdraw( )* has two possible outcomes: *OK, NO*
- **Partial operations** are operations that are not defined in all database states
- *withdraw( )* can be decomposed into two partial operations, which cover all database states:
    - *withdrawOK( )*
    - *withdrawNO( )*

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Partial Operations

- Example: account object
    - *deposit( )*: defined in all initial states (total)
    - *withdrawOK(acct,x)*: defined in all states in which $bal \geq x$ (partial)
    - *withdrawNO(acct,x)*: defined in all states in which $bal < x$ (partial)
- When a transaction submits *withdraw( ),* control checks balance and converts to either *withdrawOK( )* or *withdrawNO( )* and acquires appropriate lock

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Partial Operations

- Partial operations allow even more semantics to be introduced
- Insight: while *deposit( )* does not commute with *withdraw( )*, it does (backward) commute with *withdrawOK( )*

withdrawOK(a,n) deposit(a,m) $\rightarrow$ *deposit(a,m) withdrawOK(a.n)*

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Backward Commutativity

- *p backward commutes* through *q* iff in all states in which the sequence *q, p* is defined, the sequence *p, q* is defined and
    - *p* and *q* return the same information in both and
    - The database is left in the same final state
- Example:
    - *deposit(a,m)* backward commutes through *withdrawOK(a,n)*
        - In all database states in which *withdrawOK(a,n), deposit(a,m)* is defined, *deposit(a,m), withdrawOK(a,n)* is also defined.
    - *withdrawOK(a,n)* does not backward commute through *deposit(a,m)*
    - Backward commute is not symmetric

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A Concurrency Control Based on Partial Abstract Operations

Requested Mode	Granted Mode: *deposit( )*	*withdrawOK( )*	*withdrawNO( )*
deposit( )			X
withdrawOK( )	X		
withdrawNO( )		X	

- Control grants *deposit*, *withdrawOK*, and *withdrawNO* locks
    - Conflict relation is
        - not symmetric
        - based on backward commutativity

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A Concurrency Control Based on Partial Abstract Operations

- **Advantage**: Increased concurrency and hence increased transaction throughput
- **Disadvantage**: Concurrency control has to access the database to determine the return value (hence the operation requested) before consulting table
- Hence (with an immediate update system) if T writes *x* and later aborts, physical restoration can be used.

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Atomicity and Abstract Operations

- A write operation (the only conventional operation that modifies items) conflicts with *all* other operations on the same data
- **Physical restoration** (restore original value) does not work with abstract operations since two operations that modify a data item might commute
  - How do you handle the schedule: $\ldots p_1(x)\ q_2(x)\ abort_1 \ldots$ if both operations modify x?
- **Logical restoration** (with compensating operations) must be used
  - *e.g., increment(x)* compensates for *decrement(x)*

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A Closer Look at Compensation

- We have discussed compensation before, but now we want to use it in combination with locking to guarantee serializability and atomicity
- We must define compensation more carefully

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Requirements for an Operation to Have a Compensating Operation

- For an operation to have a compensating operation, it must be one-to-one
  - For each input there is a unique output
  - The parameters of the compensating operation are the same as the parameters of the operation being compensated
    - *increment(x)* compensate *decrement(x)*

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Logical Restoration (Compensation)

- Consider schedule: $p_1(x)\ q_2(x)\ abort_1$
- $q_2(x)$ must (backward) commute through $p_1(x)$, since the concurrency control scheduled the operation
- This is equivalent to $q_2(x)\ p_1(x)\ abort_1$
- Then $abort_1$ can be implemented with a compensating operation: $q_2(x)\ p_1(x)\ p_1^{-1}(x)$
  - This is equivalent to $q_2(x)$
- Thus $p_1(x)\ q_2(x)\ p_1^{-1}(x)$ is equivalent to $q_2(x)$

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Logical Restoration (Compensation)

- Example:

$p_1(x) = decrement(x)$
$p_1^{-1}(x) = increment(x)$

compensating operation

$$decrement_1(x)\ increment_2(x)\ increment_1(x) \equiv increment_2(x)$$

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Undo Operations

- Not all operations have compensating operations
  - For example, *reset(x)*, which sets *x* to *0*, is not one-to-one and has no compensating operation
  - It does have an undo operation, *set(x, X)*, which sets the value of *x* to what it was right before *reset(x)* was executed.

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The Previous Approach Does Not Work

$reset_1(x)\ reset_2(x)\ set_1(x, X_1)$

- Since the two *reset*s commute, we can rewrite the schedule as

  $reset_2(x)\ reset_1(x)\ set_1(x, X_1)$

- But this schedule does not undo the result of $\text{reset}_1(\text{x})$, because the value when $reset_1(x)$ starts is different in the second schedule

What to Do with Undo Operations

- One approach is to require that the operation get an exclusive lock, so that no other operation can come between an operation and its undo operation

Another Approach

- Suppose p^{undo} commutes with *q*. Then

  $p\ q\ p^{undo} \equiv p\ p^{undo}\ q$

- Now *p* has the same initial value in both schedules, and thus the undo operation works correctly.

Another Approach

- Theorem
  - Serializability and recoverability is guaranteed if the condition under which an operation *q* does not conflict with a previously granted operation *p* is
    - *q* backward commutes through *p*, and
    - Either *p* has a compensating operation, or when a *p* lock is held, p^{undo} backward commutes through *q*

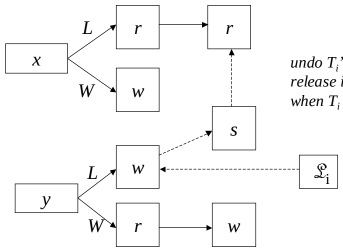
Still Another Approach

- Sometimes we can decompose an operation that does not have a compensating operation into two partial operations, each of which does have a compensating operation
  - *withdraw(x)* does not have a compensating operation
    - Depending on the initial value of the account, it might perform the withdrawal and decrement that value by x or it might just return no
    - It has an undo operation, *conditionalDeposit(x,y)*
  - The two partial operations, *withdrawOK(x)* and *withdrawNO(x)* are one-to-one and hence do have compensating operations.

Locking Implementation of Savepoints

- When T_i creates a savepoint, *s*, insert a marker for *s* in T_i's lock list, $\mathfrak{L}_i$, that separates lock entries acquired before creation from those acquired after creation
- When T_i rolls back to *s*, release all locks preceding marker for *s* in $\mathfrak{L}_i$ (in addition to undoing all updates made since savepoint creation)

Locking Implementation



undo T_i's update to y and release its write lock when T_i rolls back to s

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Locking Implementation of...

- *Chaining*: nothing new
- *Recoverable queue*: Since queue is implemented by a separate server (different from DBMS), the locking discipline need not be two-phase; discipline can be designed to suit the semantics of (the abstract operations) *enqueue* and *dequeue*
  - Lock on head (tail) pointer released when dequeue (enqueue) operations complete
    - Hence not strict or isolated
  - Lock on entry that is enqueued or dequeued held to commit time

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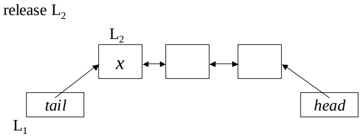
Recoverable Queue

begin transaction
….
enqueue(x) ← acquire L_1, L_2
← release L_1
….
commit ← release L_2



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Locking Implementation of Nested Transactions

- Nested transactions satisfy:
  - Nested transactions are isolated with respect to one another
  - A parent does not execute concurrently with its children
  - A child (and its descendants) is isolated from its siblings (and their descendants)

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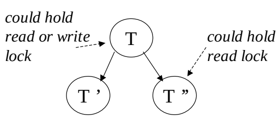
Locking Implementation of Nested Transactions

- A request to read *x* by subtransaction T′ of nested transaction T is granted if:
  - No other nested transaction holds a write lock on *x*
  - All other subtransactions of T holding write locks on *x* are ancestors of T′ (hence are not executing)



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Intuition

- A request to read *x* by subtransaction T' of nested transaction T is granted even though an ancestor of T' holds a write lock on *x*

without nesting		with nesting
T: begin transaction		*T: begin transaction*
…		…
w(x)		*w(x)*
		T': begin transaction
…		…
r(x)	*r(x) does not conflict with w(x)*	*r(x)*
…		…
		commit
commit		*commit*

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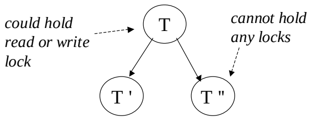
Locking Implementation of Nested Transactions

- A request to write *x* by subtransaction T ' of nested transaction T is granted if:
  - No other nested transaction holds a read or write lock on *x*
  - All other subtransactions of T holding read or write locks on *x* are ancestors of T ' (and hence are not executing)



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Locking Implementation of Nested Transactions

- All locks obtained by T' are held until it completes
  - If it aborts, all locks are discarded
  - If it commits, any locks it holds that are not held by its parent are inherited by its parent
- When top-level transaction (and hence entire nested transaction) commits, all locks are discarded

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Locking Implementation of Multilevel Transactions

- Generalization of strict two-phase locking concurrency control
  - Uses semantics of operations at each level to determine commutativity
  - Uses different concurrency control at each level

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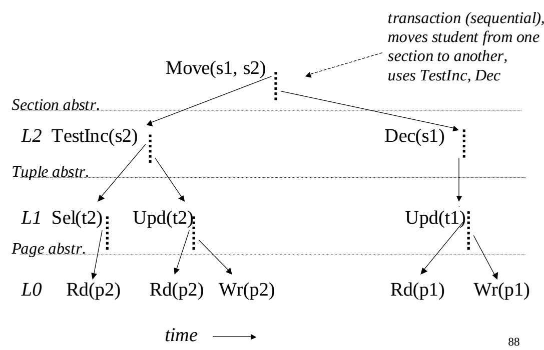
Example - Switch Sections



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Multilevel Transactions

- **Example**:
  - *Move(s1,s2)* produces *TestInc(s2), Dec(s1)*
  - *$Move_1$(s1,s2), $Move_2$(s1, s3)* might produce
    $TestInc_1$(s2), $TestInc_2$(s3), Dec_2(s1), Dec_1(s1)
  - Since two *Dec* operations on the same object commute (they do not impose an ordering among transactions), this schedule is equivalent to
    $TestInc_1$(s2), Dec_1(s1), $TestInc_2$(s3), Dec_2(s1)
    and hence could be allowed by a multilevel control, but ...

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Multilevel Control

- **Problem**: A control assumes that the execution of operations it schedules is isolated: If op_1 and op_2 do not conflict, they can be executed concurrently and the result will be either op_1, op_2 or op_2, op_1
  - Not true in a multilevel control where an operation is implemented as a program at the next lower level that might invoke multiple operations at the level below. Hence, concurrent operations at one level might not be totally ordered at the next

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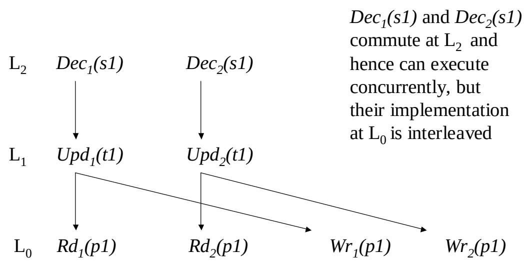
Multilevel Transactions



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Guaranteeing Operation Isolation

- **Solution**: Use a concurrency control at each level
  - L_i receives a request from L_{i+1} to execute *op*
  - Concurrency control at L_i, CC_i, schedules *op* to be executed; it assumes execution is isolated
  - *op* is implemented as a program, *P*, in L_i
  - *P* is executed as a subtransaction so that it is serializable with respect to other operations scheduled by CC_i
  - Serializability guaranteed by CC_{i-1}

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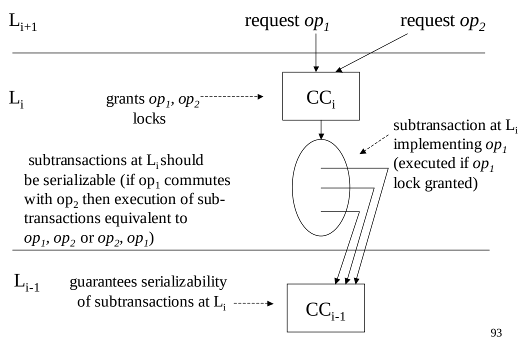
Guaranteeing Operation Isolation



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A Multilevel Concurrency Control for the Example

- The control at L_2 uses *TestInc* and *Dec* locks
- The control at L_1 uses *Sel* and *Upd* locks
- The control at L_0 uses *Rd* and *Wr* locks

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Timestamp-Ordered Concurrency Control

- Each transaction given a (unique) timestamp (current clock value) when initiated
- Uses the immediate update model
- Guarantees equivalent serial order based on timestamps (initiation order)
  - Control is *static* (as opposed to *dynamic*, in which the equivalent serial order is determined as the schedule progresses)

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Timestamp-Ordered Concurrency Control

- Associated with each database item, *x*, are two timestamps:
  - *wt(x)*, the largest timestamp of any transaction that has written *x*,
  - *rt(x)*, the largest timestamp of any transaction that has read *x*,
  - and an indication of whether or not the last write to that item is from a committed transaction

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Timestamp-Ordered Concurrency Control

- If T requests to read *x*:
  - **R1**: if $TS(T) < wt(x)$, then T is too old; abort T
  - **R2**: if $TS(T) > wt(x)$, then
    - if the value of *x* is committed, grant T's read and if $TS(T) > rt(x)$ assign $TS(T)$ to $rt(x)$
    - if the value of *x* is not committed, T waits (to avoid a dirty read)

Timestamp-Ordered Concurrency Control

- If T requests to write *x* :
  - **W1**: If $TS(T) < rt(x)$, then T is too old; abort T
  - **W2**: If $rt(x) < TS(T) < wt(x)$, then no transaction that read *x* should have read the value T is attempting to write and no transaction will read that value (See R1)
    - If *x* is committed, grant the request but do not do the write
      - This is called the Thomas Write Rule
    - If *x* is not committed, T waits to see if newer value will commit. If it does, discard T's write, else perform it
  - **W3**: If $wt(x), rt(x) < TS(T)$, then if *x* is committed, grant the request and assign $TS(T)$ to $wt(x)$, else T waits

Example

- Assume $TS(T_1) < TS(T_2)$, at t_0 *x* and *y* are committed, and *x*'s and *y*'s read and write timestamps are less than $TS(T_1)$

T_1 :		*r(y)*				*w(x)*	*commit*
T_2:			*w(y)*	*w(x)*	*commit*		
	t_0	t_1	t_2	t_3		t_4	

t_1: (R2) $TS(T_1) > wt(y)$; assign $TS(T_1)$ to $rt(y)$
t_2: (W3) $TS(T2) > rt(y), wt(y)$; assign $TS(T_2)$ to $wt(y)$
t_3: (W3) $TS(T2) > rt(x), wt(x)$; assign $TS(T_2)$ to $wt(x)$
t_4: (W2) $rt(x) < TS(T1) < wt(x)$; grant request, but do not do the write

Timestamp-Ordered Concurrency Control

- Control accepts schedules that are *not conflict equivalent* to any serial schedule and would not be accepted by a two-phase locking control
  - Previous example equivalent to T_1, T_2
- But additional space required in database for storing timestamps and time for managing timestamps
  - Reading a data item now implies writing back a new value of its timestamp

Optimistic Algorithms

- Do task under simplifying (optimistic) assumption
  - **Example**: Operations rarely conflict
- Check afterwards if assumption was true.
  - **Example**: Did a conflict occur?
- Redo task if assumption was false
  - **Example**: If a conflict has occurred rollback, else commit
- Performance benefit if assumption is generally true and check can be done efficiently

Optimistic Concurrency Control

- *Under the optimistic assumption that conflicts do not occur*, read and write requests are always granted (no locking, no overhead!)
- *Since conflicts might occur*:
  - Database might be corrupted if writes were immediate, hence a deferred-update model is used
  - Transaction has to be "validated" when it completes
    - If a conflict has occurred abort (but no rollback is necessary) and redo transaction
- Approach contrasts with pessimistic control which assumes conflicts are likely, takes preventative measures (locking), and does no validation

Optimistic Concurrency Control

- Transaction has three phases:
  - Begin transaction
    - *Read Phase* - transaction executes: reads from database, writes to intentions list (deferred-update, no changes to database)
  - Request commit
    - *Validation Phase* - check whether conflicts occurred during read phase; if yes abort (discard intentions list)
  - Commit
    - *Write Phase* - write intentions list to database (deferred update) if validation successful
  - For simplicity, we assume here that validation and write phases form a single critical section (only one transaction is in its validation/write phase at a time)

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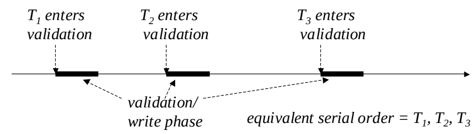
Optimistic Concurrency Control

- Guarantees an equivalent serial schedule in which the order of transactions is the order in which they enter validation (dynamic)
- For simplicity, we will assume that validation and write phases form a single critical section (only one transaction is in its validation/write phase at a time)



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Validation

- When T_1 enters validation, a check is made to see if T_1 conflicted with any transaction, T_2, that entered validation at an earlier time
- Check uses two sets constructed during read phase:
  - $R(T_1)$: identity of all database items T_1 read
  - $W(T_1)$: identity of all database items T_1 wrote

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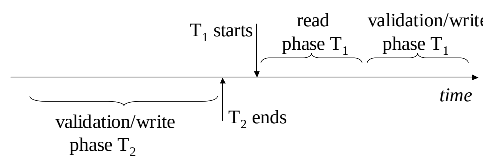
Validation

- *Case 1*: T_1's read phase started after T_2 finished its validation/write phase
  - T_1 follows T_2 in all conflicts, hence commit T_1 (T_1 follows T_2 in equivalent serial order)



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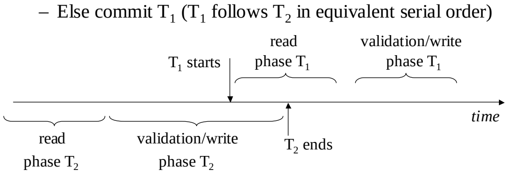
Validation

- *Case 2*: T_1's read phase overlaps T_2's validation/write phase
  - If $WS(T_2) \cap RS(T_1) \neq \Phi$, then abort T_1
    - A read of T_1 *might* have preceded a write of T_2 – a *possible* violation of equivalent serial order
  - Else commit T_1 (T_1 follows T_2 in equivalent serial order)



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Validation

- *Case 3*: T_1's validation/write phase overlaps T_2's validation/write phase
  - Cannot happen since we have assumed that validation/write phases do not overlap
- Hence, all possible overlaps of T_1 and T_2 have been considered

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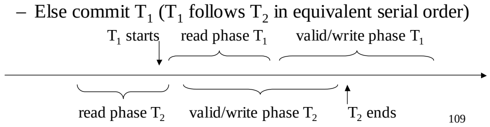
Validation

- A more practical optimistic control allows case 3 and avoids the bottleneck implied by only allowing only one transaction at a time in the validation/write phase.
- *Case 3*: T_1's validation/write phase overlaps T_2's validation/write phase
  - If $WS(T_2) \cap (WS(T_1) \cup RS(T_1)) \neq \Phi$, then abort T_1
    - A read *or* write of T_1 *might* have preceded a write of T_2 – a violation of equivalent serial order
  - Else commit T_1 (T_1 follows T_2 in equivalent serial order)



Optimistic Concurrency Control

- No locking (and hence no waiting) means deadlocks are not possible
- Rollback is a problem if optimistic assumption is not valid: work of entire transaction is lost
  - With two-phase locking, rollback occurs only with deadlock
  - With timestamp-ordered control, rollback is detected before transaction completes