CSE 544
Principles of Database Management Systems

Lecture 14 and 15 –
Transactions: Concurrency Control

Announcmenets

- Last homework due on Friday
- Last reading assignment next Monday
- Projects
  - Milestones: will send comments by email
  - Posters/demos: next Tuesday, March 6, 10am – 2pm

References

- **Database management systems.**
  Ramakrishnan and Gehrke.
  Third Ed. **Chapters 16 and 17.**

Outline

- Transactions motivation, definition, properties
- Concurrency control and locking
- Optimistic concurrency control

Motivating Example

```
UPDATE Budget
SET money=money-100
WHERE pid = 1

UPDATE Budget
SET money=money+60
WHERE pid = 2

UPDATE Budget
SET money=money+40
WHERE pid = 3
```

```
SELECT sum(money)
FROM Budget
```

Would like to treat each group of instructions as a unit

Definition

A transaction = one or more operations, single real-world transition

Examples

- Transfer money between accounts
- Purchase a group of products
- Register for a class (either waitlist or allocated)
- ...

Transactions

- Major component of database systems
- Critical for most applications; arguably more so than SQL

- Fact: Turing awards to database researchers:
  - Charles Bachman 1973 for CODASYL
  - Edgar Codd 1981 for relational model
  - Jim Gray 1998 for transactions
  - Michael Stonebraker 2015 for postgres

Transaction Example

```
START TRANSACTION
    UPDATE Budget
        SET money = money - 100
        WHERE pid = 1
    UPDATE Budget
        SET money = money + 60
        WHERE pid = 2
    UPDATE Budget
        SET money = money + 40
        WHERE pid = 3
COMMIT
```

ROLLBACK

- If the application gets to a place where it can’t complete the transaction successfully, it can execute **ROLLBACK**
- This causes the system to “abort” the transaction
- Database returns to a state without any of the changes made by the transaction

Reasons for Rollback

- User changes their mind (“ctl-C”/cancel)
- Explicit in program, when app program finds a problem
  - e.g., when qty on hand < qty being sold
- System-initiated abort
  - System crash
  - Housekeeping, e.g., due to timeouts, admission control, etc

ACID Properties

- Atomicity: Either all changes performed by transaction occur or none occurs
- Consistency: A transaction as a whole does not violate integrity constraints
- Isolation: Transactions appear to execute one after the other in sequence
- Durability: If a transaction commits, its changes will survive failures

What Could Go Wrong?

- Why is it hard to provide ACID properties?

- Concurrent operations
  - Isolation problems
  - We saw one example earlier

- Failures can occur at any time
  - Atomicity and durability problems
  - Next week

- Transaction may need to abort

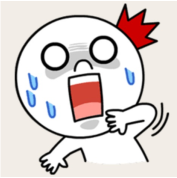
What Could Go Wrong

```
Client 1: INSERT INTO SmallProduct(name, price)
          SELECT pname, price
          FROM Product
          WHERE price <= 0.99

          DELETE Product
          WHERE price <=0.99

Client 2: SELECT count(*)
          FROM Product

          SELECT count(*)
          FROM SmallProduct
```



Inconsistent reads

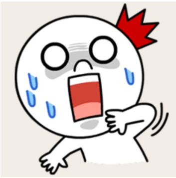
What Could Go Wrong

Client 1:

```
UPDATE Product
SET Price = Price – 1.99
WHERE pname = ‘Gizmo’
```

Client 2:

```
UPDATE Product
SET Price = Price*0.5
WHERE pname=‘Gizmo’
```



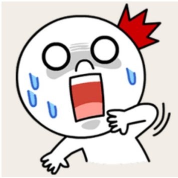
Lost update

What Could Go Wrong

```
Client 1: UPDATE Account
                    SET amount = 1000000
                    WHERE number = 1001
```

Aborted by system

```
Client 2: SELECT Account.amount
              FROM Account
              WHERE Account.number = 1001
```



Dirty reads

Summary of What Can Go Wrong

- Concurrent execution problems
  - Write-read conflict: dirty read (includes inconsistent read)
    - A transaction reads a value written by another transaction that has not yet committed
  - Read-write conflict: unrepeatable read
    - A transaction reads the value of the same object twice. Another transaction modifies that value in between the two reads
  - Write-write conflict: lost update
    - Two transactions update the value of the same object. The second one to write the value overwrite the first change
- Failure problems
  - DBMS can crash in the middle of a series of updates
  - Can leave the database in an inconsistent state

ACID Properties

- Atomicity: Either all changes performed by transaction occur or none occurs
- Consistency: A transaction as a whole does not violate integrity constraints
- Isolation: Transactions appear to execute one after the other in sequence
- Durability: If a transaction commits, its changes will survive failures

Outline

- Transactions motivation, definition, properties
- Concurrency control and locking
- Optimistic concurrency control

Schedules

- Given multiple transactions

- A *schedule* is a sequence of interleaved actions from all transactions

Example: Two Transactions

T1	T2
READ(A, t)	READ(A, s)
t := t+100	s := s*2
WRITE(A, t)	WRITE(A,s)
READ(B, t)	READ(B,s)
t := t+100	s := s*2
WRITE(B,t)	WRITE(B,s)

A Serial Schedule

Time	T1	T2
	READ(A, t)	
	t := t+100	
	WRITE(A, t)	
	READ(B, t)	
	t := t+100	
	WRITE(B,t)	
		READ(A,s)
		s := s*2
		WRITE(A,s)
		READ(B,s)
		s := s*2
		WRITE(B,s)

Serializable Schedule

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

A Serializable Schedule

T1	T2
READ(A, t)	
t := t+100	
WRITE(A, t)	
	READ(A,s)
	s := s*2
	WRITE(A,s)
READ(B, t)	
t := t+100	
WRITE(B,t)	
	READ(B,s)
	s := s*2
	WRITE(B,s)

Notice:
This is NOT a serial schedule

A Non-Serializable Schedule

T1	T2
READ(A, t)	
t := t+100	
WRITE(A, t)	
	READ(A,s)
	s := s*2
	WRITE(A,s)
	READ(B,s)
	s := s*2
	WRITE(B,s)
READ(B, t)	
t := t+100	
WRITE(B,t)	

Checking Serializability

- Goal: build a <u>*scheduler*</u> that guarantees serializability
- But how do we know that a schedule is serializable?
  - In general, this is undecidable:
    E.g. T1, T2 compute complex functions, do they commute?
- Two simple sufficient (but not necessary) conditions:
  - Conflict serializability
  - View serializability

Ignoring Details

Capture only the read/write actions
Ignore the computations (assume worse case)

T_1: $r_1(A)$; $w_1(A)$; $r_1(B)$; $w_1(B)$
T_2: $r_2(A)$; $w_2(A)$; $r_2(B)$; $w_2(B)$

Key Idea: Focus on *conflicting* operations

Conflict Serializability

Conflicts: (i.e., swapping will change program behavior)

Two actions by same transaction T_i: $\boxed{r_i(X); w_i(Y)}$

Two writes by T_i, T_j to same element $\boxed{w_i(X); w_j(X)}$

Read/write by T_i, T_j to same element $\boxed{w_i(X); r_j(X)}$

$\boxed{r_i(X); w_j(X)}$

Conflict Serializability

- A schedule is *conflict serializable* if it can be transformed into a serial schedule by a series of swappings of adjacent non-conflicting actions

- Every conflict-serializable schedule is serializable
- The converse is not true (why?)

Conflict Serializability

Example:

$r_1(A); w_1(A); r_2(A); w_2(A); r_1(B); w_1(B); r_2(B); w_2(B)$

Conflict Serializability

Example:

$r_1(A); w_1(A); r_2(A); w_2(A); r_1(B); w_1(B); r_2(B); w_2(B)$

$\Downarrow$

$r_1(A); w_1(A); r_1(B); w_1(B); r_2(A); w_2(A); r_2(B); w_2(B)$

Conflict Serializability

Example:

$r_1(A); w_1(A); r_2(A); \boxed{w_2(A); r_1(B);} w_1(B); r_2(B); w_2(B)$

$\Downarrow$

$r_1(A); w_1(A); r_1(B); w_1(B); r_2(A); w_2(A); r_2(B); w_2(B)$

Conflict Serializability

Example:

$r_1(A); w_1(A); r_2(A); \boxed{w_2(A); r_1(B);} w_1(B); r_2(B); w_2(B)$

$r_1(A); w_1(A); \boxed{r_2(A); r_1(B);} w_2(A); w_1(B); r_2(B); w_2(B)$

$r_1(A); w_1(A); r_1(B); w_1(B); r_2(A); w_2(A); r_2(B); w_2(B)$

Conflict Serializability

Example:

$r_1(A); w_1(A); r_2(A); \boxed{w_2(A); r_1(B);} w_1(B); r_2(B); w_2(B)$

$r_1(A); w_1(A); \boxed{r_2(A); r_1(B);} w_2(A); w_1(B); r_2(B); w_2(B)$

$r_1(A); w_1(A); r_1(B); r_2(A); \boxed{w_2(A); w_1(B)}; r_2(B); w_2(B)$

….

$r_1(A); w_1(A); r_1(B); w_1(B); r_2(A); w_2(A); r_2(B); w_2(B)$

Testing for Conflict-Serializability

Precedence graph:

- A node for each transaction T_i,
- An edge from T_i to T_j whenever an action in T_i conflicts with, and comes before an action in T_j

- The schedule is conflict-serializable iff the precedence graph is acyclic

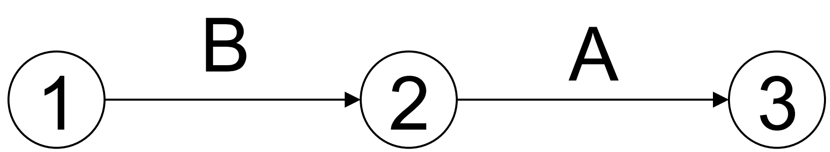
Example 1

$r_2(A); r_1(B); w_2(A); r_3(A); w_1(B); w_3(A); r_2(B); w_2(B)$

1 2 3

Example 1

$r_2(A); r_1(B); w_2(A); r_3(A); w_1(B); w_3(A); r_2(B); w_2(B)$



This schedule is conflict-serializable

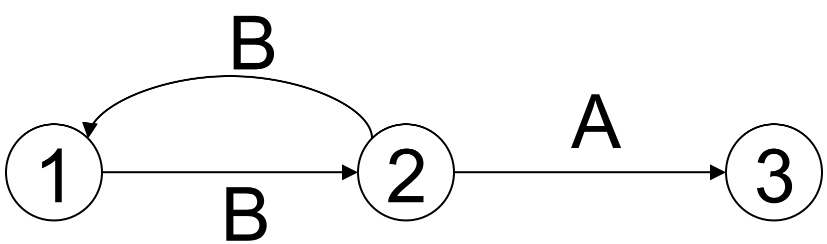
Example 2

$r_2(A); r_1(B); w_2(A); r_2(B); r_3(A); w_1(B); w_3(A); w_2(B)$

1 2 3

Example 2

$r_2(A); r_1(B); w_2(A); r_2(B); r_3(A); w_1(B); w_3(A); w_2(B)$



This schedule is NOT conflict-serializable

View Equivalence

- A serializable schedule need not be conflict serializable, even under the “worst case update” assumption

$w_1(X); w_2(X); w_2(Y); w_1(Y); w_3(Y);$

Is this schedule conflict-serializable ?

View Equivalence

- A serializable schedule need not be conflict serializable, even under the “worst case update” assumption

$w_1(X); w_2(X); w_2(Y); w_1(Y); w_3(Y);$

Is this schedule conflict-serializable ?

No…

View Equivalence

- A serializable schedule need not be conflict serializable, even under the "worst case update" assumption

$w_1(X); w_2(X); w_2(Y); w_1(Y); w_3(Y);$

Lost write

$w_1(X); w_1(Y); w_2(X); w_2(Y); w_3(Y);$

Equivalent, but not conflict-equivalent

View Equivalence

T1	T2	T3
W1(X)		
	W2(X)	
	W2(Y)	
	CO2	
W1(Y)		
CO1		
	Lost	W3(Y)
		CO3

T1	T2	T3
W1(X)		
W1(Y)		
CO1		
	W2(X)	
	W2(Y)	
	CO2	
		W3(Y)
		CO3

Serializable, but not conflict serializable

Scheduler

- The scheduler is the module that schedules the transaction’s actions, ensuring serializability
- How? We discuss three techniques in class:
  - Locks
  - Timestamps
  - Validation

Outline

- Transactions motivation, definition, properties
- Concurrency control and locking
- Optimistic concurrency control

Locking Scheduler

Simple idea:

- Each element has a unique lock
- Each transaction must first acquire the lock before reading/writing that element
- If lock is taken by another transaction, then wait
- The transaction must release the lock(s)

Notation

$l_i(A)$ = transaction T_i acquires lock for element A

$u_i(A)$ = transaction T_i releases lock for element A

Example

T1	T2
$L_1(A)$; READ(A, t)	
t := t+100	
WRITE(A, t); $U_1(A)$; $L_1(B)$	
	$L_2(A)$; READ(A,s)
	s := s*2
	WRITE(A,s); $U_2(A)$;
	$L_2(B)$; **DENIED…**
READ(B, t)	
t := t+100	
WRITE(B,t); $U_1(B)$;	
	…**GRANTED;** READ(B,s)
	s := s*2
	WRITE(B,s); $U_2(B)$;

Scheduler has ensured a conflict-serializable schedule

Is this enough?

T1	T2
$L_1(A)$; READ(A, t)	
t := t+100	
WRITE(A, t); $U_1(A)$;	
	$L_2(A)$; READ(A,s)
	s := s*2
	WRITE(A,s); $U_2(A)$;
	$L_2(B)$; READ(B,s)
	s := s*2
	WRITE(B,s); $U_2(B)$;
$L_1(B)$; READ(B, t)	
t := t+100	
WRITE(B,t); $U_1(B)$;	

Locks did not enforce conflict-serializability !!!

Two Phase Locking (2PL)

The 2PL rule:

- In every transaction, all lock requests must preceed all unlock requests

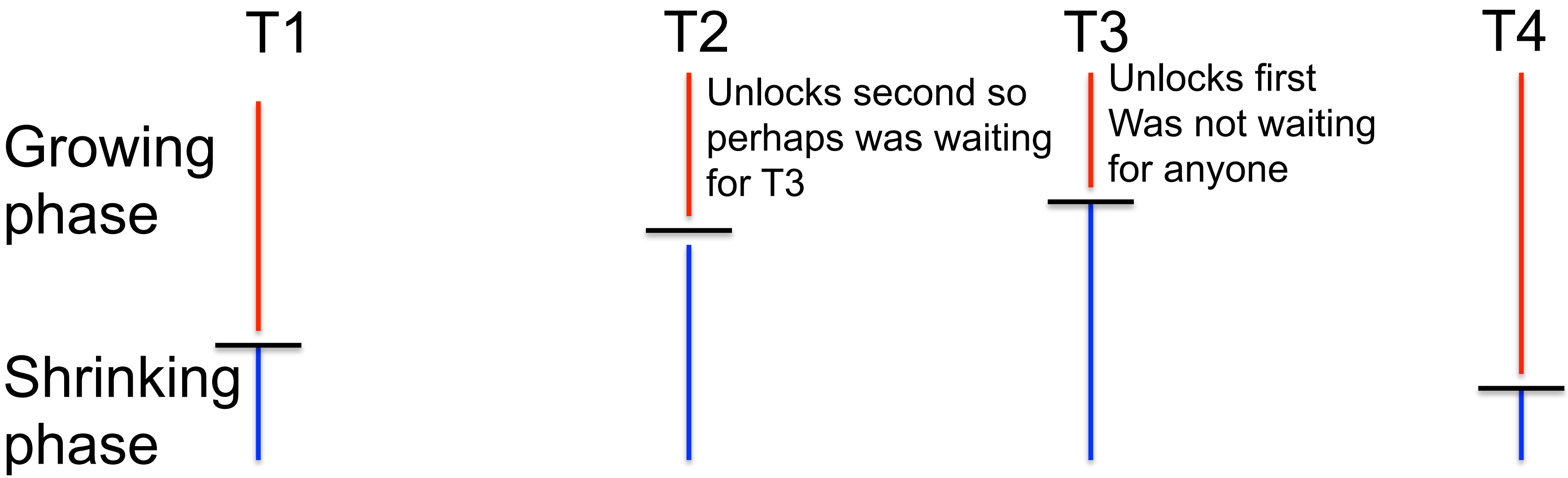
- This ensures conflict serializability ! (why?)

Example: 2PL transactions

T1	T2
$L_1(A)$; $L_1(B)$; READ(A, t)	
t := t+100	
WRITE(A, t); $U_1(A)$	
	$L_2(A)$; READ(A,s)
	s := s*2
	WRITE(A,s);
	$L_2(B)$; **DENIED…**
READ(B, t)	
t := t+100	
WRITE(B,t); $U_1(B)$;	
	…**GRANTED;** READ(B,s)
	s := s*2
	WRITE(B,s); $U_2(A)$; $U_2(B)$;

Now it is conflict-serializable

Example with Multiple Transactions



Equivalent to each transaction executing entirely the moment it enters shrinking phase

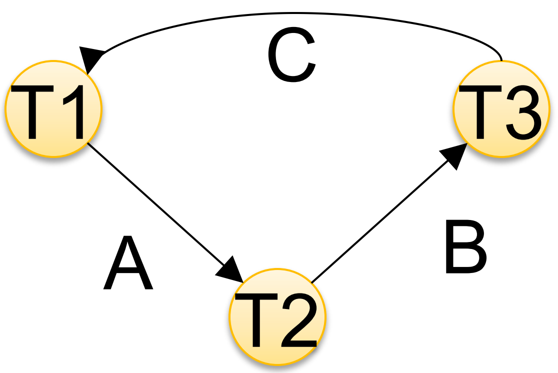
Two Phase Locking (2PL)

Theorem: 2PL ensures conflict serializability

Two Phase Locking (2PL)

Theorem: 2PL ensures conflict serializability

Proof. Suppose not: then there exists a cycle in the precedence graph.



Two Phase Locking (2PL)

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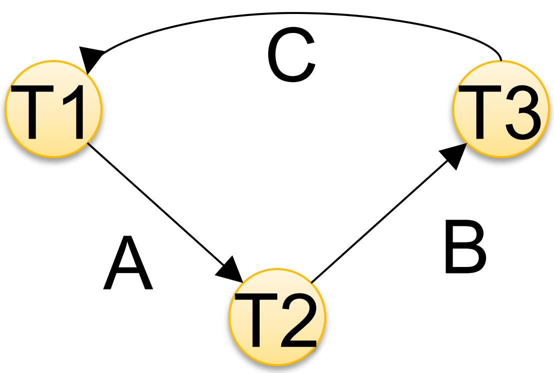

Then there is the following **temporal** cycle in the schedule:

Two Phase Locking (2PL)

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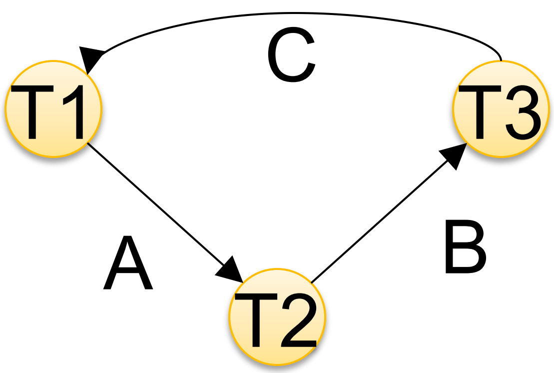
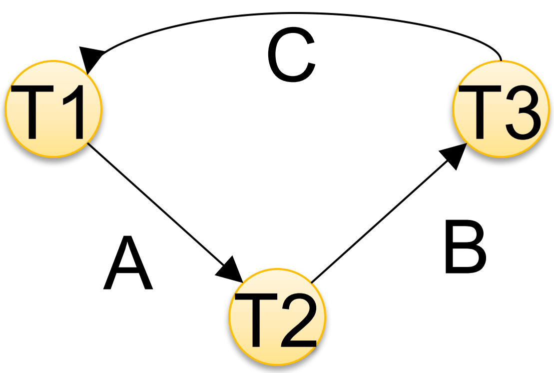


Then there is the following **temporal** cycle in the schedule:
$U_1(A) \rightarrow L_2(A)$ why?

Two Phase Locking (2PL)

Theorem: 2PL ensures conflict serializability

Proof. Suppose not: then there exists a cycle in the precedence graph.



Then there is the following <u>**temporal**</u> cycle in the schedule:

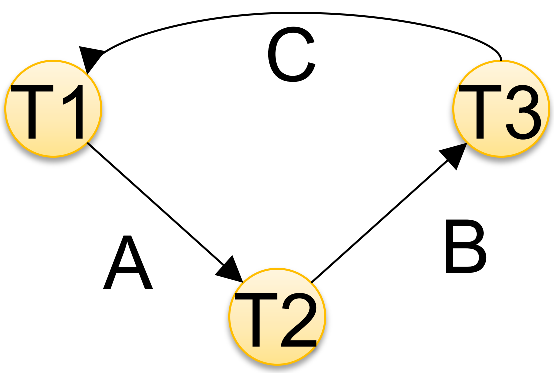
$U_1(A) \rightarrow L_2(A)$

$L_2(A) \rightarrow U_2(B)$ why?

Two Phase Locking (2PL)

Theorem: 2PL ensures conflict serializability

Proof. Suppose not: then there exists a cycle in the precedence graph.



Then there is the following ***temporal*** cycle in the schedule:

$U_1(A) \rightarrow L_2(A)$
$L_2(A) \rightarrow U_2(B)$
$U_2(B) \rightarrow L_3(B)$
$L_3(B) \rightarrow U_3(C)$
$U_3(C) \rightarrow L_1(C)$
$L_1(C) \rightarrow U_1(A)$

Contradiction

A New Problem: Non-recoverable Schedule

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A :=A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B :=B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;
	Commit
Rollback	

A New Problem: Non-recoverable Schedule

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A :=A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B :=B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;
	Commit
Rollback	

Elements A, B written by T1 are restored to their original value.

A New Problem: Non-recoverable Schedule

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A :=A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B :=B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;
	Commit
Rollback	

Dirty reads of A, B lead to incorrect writes.

Elements A, B written by T1 are restored to their original value.

A New Problem: Non-recoverable Schedule

T1	T2
$L_1(A)$; $L_1(B)$; READ(A)	
A :=A+100	
WRITE(A); $U_1(A)$	
	$L_2(A)$; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; BLOCKED…
READ(B)	
B :=B+100	
WRITE(B); $U_1(B)$;	
	…GRANTED; READ(B)
	B := B*2
	WRITE(B); $U_2(A)$; $U_2(B)$;
	Commit
Rollback	

Dirty reads of A, B lead to incorrect writes.

Elements A, B written by T1 are restored to their original value.

Can no longer undo!

Strict 2PL

The Strict 2PL rule:

All locks are held until commit/abort:
All unlocks are done *together* with commit/abort.

With strict 2PL, we will get schedules that are both conflict-serializable and recoverable

Strict 2PL

T1	T2
$L_1(A)$; READ(A)	
A :=A+100	
WRITE(A);	
	$L_2(A)$; BLOCKED…
$L_1(B)$; READ(B)	
B :=B+100	
WRITE(B);	
Rollback & $U_1(A)$;$U_1(B)$;	
	…GRANTED; READ(A)
	A := A*2
	WRITE(A);
	$L_2(B)$; READ(B)
	B := B*2
	WRITE(B);
	Commit & $U_2(A)$; $U_2(B)$;

Strict 2PL

- Lock-based systems always use strict 2PL
- Easy to implement:
  - Before a transaction reads or writes an element A, insert an L(A)
  - When the transaction commits/aborts, then release all locks
- Ensures both conflict serializability and recoverability

Deadlock

- Transaction T_1 waits for a lock held by T_2;
- T_2 waits for T_3;
- T_3 waits for T_4;
- . . .
- T_n waits for T_1

- A deadlock is when two or more transactions are waiting for each other to complete

Handling Deadlock

- **Deadlock avoidance**
  - Acquire locks in pre-defined order
  - Acquire all locks at once before starting

- **Deadlock detection**
  - Timeouts (but hard to pick the right threshold)
  - Wait-for graph; this is what commercial systems use (they check graph periodically)

Lock Modes

- S = shared lock (for READ)
- X = exclusive lock (for WRITE)

Lock compatibility matrix:

	None	S	X
None	OK	OK	OK
S	OK	OK	Conflict
X	OK	Conflict	Conflict

Others:
U = update lock: Initially like S, later may be upgraded to X
I = increment lock (for A := A + something): Increment operations commute

Lock Granularity

- Fine granularity locking (e.g., tuples)
  - High concurrency
  - High overhead in managing locks
- Coarse grain locking (e.g., tables)
  - Many false conflicts
  - Less overhead in managing locks
- Alternative techniques
  - Hierarchical locking (and intentional locks) [commercial DBMSs]
  - Lock escalation

The Tree Protocol

- An alternative to 2PL, for tree structures
- E.g. B+ trees (the indexes of choice in databases)

- Because
  - Indexes are hot spots!
  - 2PL would lead to huge lock contention for the root node

  - Also, unlike data, the index is not directly visible to transactions
  - So only need to guarantee that index returns correct values

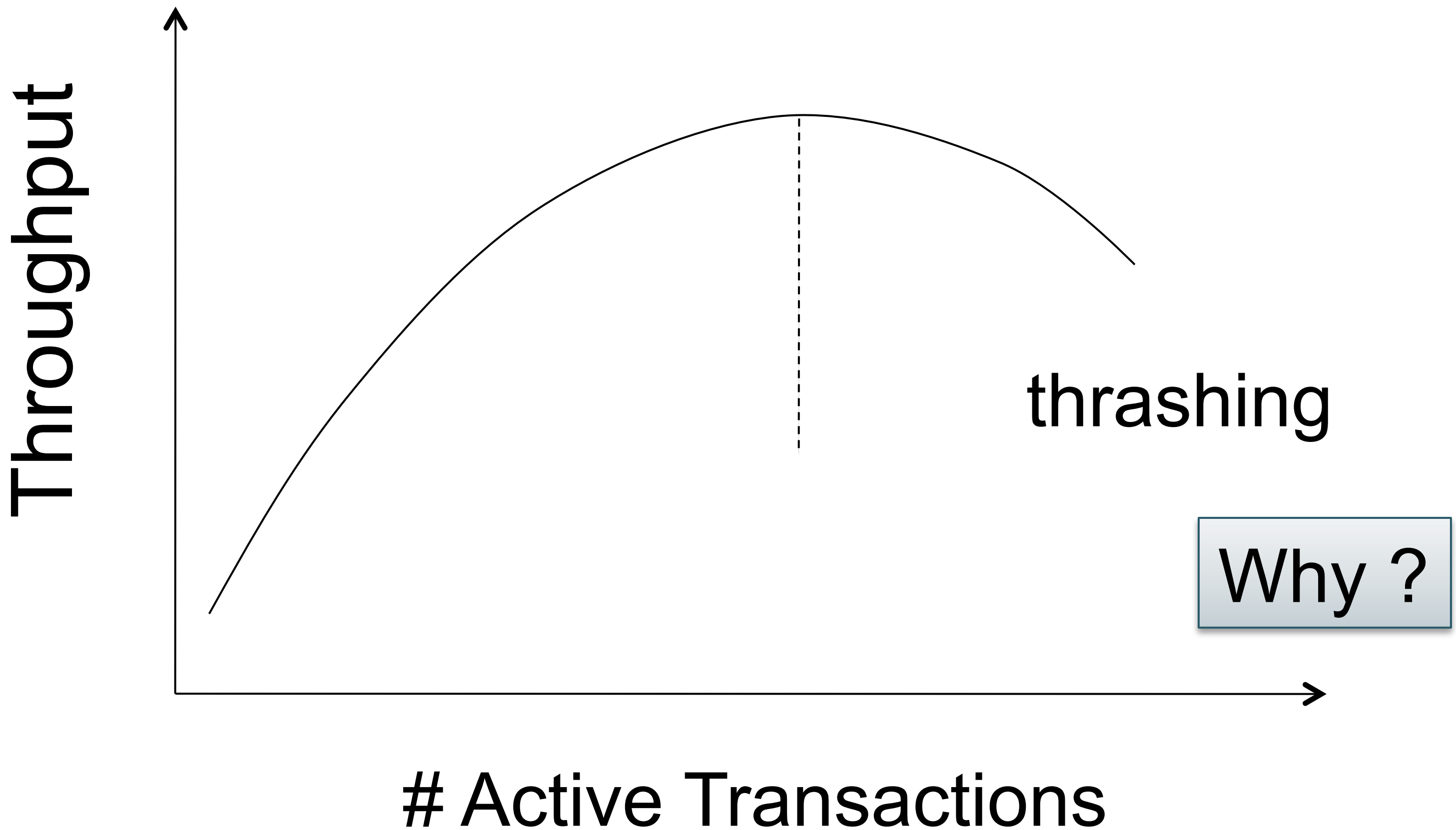
The Tree Protocol

Rules:

- A lock on a node A may only be acquired if TXN holds a lock on its parent B
- Nodes can be unlocked in any order (no 2PL necessary)
- Cannot relock a node for which already released a lock
- "Crabbing"
  - First lock parent then lock child
  - Keep parent locked only if may need to update it
  - Release lock on parent if child is not full

- The tree protocol is NOT 2PL, yet ensures conflict-serializability !
- (More in the textbook)

Lock Performance



Phantom Problem

- Static database = a fixed collection of elements (records or blocks)
  - So far we considered serializability only for a static database

- Dynamic database = elements may be inserted/deleted
  - New problem: phantoms

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color=‘blue’	
	INSERT INTO Product(name, color) VALUES (‘gizmo’,’blue’)
SELECT * FROM Product WHERE color=‘blue’	

Is this schedule serializable ?

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color='blue'	
	INSERT INTO Product(name, color) VALUES ('gizmo','blue')
SELECT * FROM Product WHERE color='blue'	

Suppose there are two blue products, X1, X2:

R1(X1),R1(X2),W2(X3),R1(X1),R1(X2),R1(X3)

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color=‘blue’	
	INSERT INTO Product(name, color) VALUES (‘gizmo’,’blue’)
SELECT * FROM Product WHERE color=‘blue’	

Suppose there are two blue products, X1, X2:

R1(X1),R1(X2),W2(X3),R1(X1),R1(X2),R1(X3)

This is conflict serializable ! What’s wrong ??

Phantom Problem

T1	T2
SELECT * FROM Product WHERE color='blue'	
	INSERT INTO Product(name, color) VALUES ('gizmo','blue')
SELECT * FROM Product WHERE color='blue'	

Suppose there are two blue products, X1, X2:

R1(X1),R1(X2),W2(X3),R1(X1),R1(X2),R1(X3)

Not serializable due to ***phantoms***

Phantom Problem

- A “phantom” is a tuple that is invisible during part of a transaction execution but not invisible during the entire execution

- In our example:
  - T1: reads list of products
  - T2: inserts a new product
  - T1: re-reads: a new product appears !

Phantom Problem

- In a ***static*** database:
  Conflict serializability
  implies view serializability
  implies serializability

- In a ***dynamic*** database, this may fail due to phantoms

- Strict 2PL guarantees conflict serializability, but not serializability

Dealing With Phantoms

Is expensive!!

- Lock the entire table, or
- Lock the index entry for ‘blue’
  - If index is available
- Or use predicate locks
  - A lock on an arbitrary predicate

Degrees of Isolation

- Isolation level “serializable” (i.e. ACID)
  - Golden standard
  - Requires strict 2PL and predicate locking
  - But often too inefficient
  - Imagine there are only a few update operations and many long read operations
- Weaker isolation levels
  - Sacrifice correctness for efficiency
  - Often used in practice (often **default**)
  - Sometimes are hard to understand

Isolation Levels in SQL

1. "Dirty reads"
   SET TRANSACTION ISOLATION LEVEL READ UNCOMMITTED

2. "Committed reads"
   SET TRANSACTION ISOLATION LEVEL READ COMMITTED

3. "Repeatable reads"
   SET TRANSACTION ISOLATION LEVEL REPEATABLE READ

4. Serializable transactions
   SET TRANSACTION ISOLATION LEVEL SERIALIZABLE

ACID

1. Isolation Level: Dirty Reads

- "Long duration" WRITE locks
  - Strict 2PL
- No READ locks
  - Read-only transactions are never delayed

Possible pbs: dirty and inconsistent reads

2. Isolation Level: Read Committed

- "Long duration" WRITE locks
  - Strict 2PL
- "Short duration" READ locks
  - Only acquire lock while reading (not 2PL)

Unrepeatable reads
When reading same element twice,
may get two different values

3. Isolation Level: Repeatable Read

- "Long duration" WRITE locks
  – Strict 2PL
- "Long duration" READ locks
  – Strict 2PL

This is not serializable yet !!!

Why ?

4. Isolation Level Serializable

- "Long duration" WRITE locks
  - Strict 2PL
- "Long duration" READ locks
  - Strict 2PL

- Deals with phantoms too

Outline

- Transactions motivation, definition, properties
- Concurrency control and locking
- Optimistic concurrency control

Locking vs Optimistic

- Locking prevents unserializable behavior from occurring: it causes transactions to wait for locks
- Optimistic methods assume no unserializable behavior will occur: they abort transactions if it does
- Locking typically better in case of high levels of contention; optimistic better otherwise

Timestamps

- Each transaction receives a unique timestamp TS(T)

Could be:

- The system's clock
- A unique counter, incremented by the scheduler

Timestamps

Main invariant:

The timestamp order defines
the serialization order of the transaction

Will generate a schedule that is view-equivalent to a serial schedule, and recoverable

Main Idea

- Scheduler receives a request, $r_T(X)$ or $w_T(X)$
- Should it allow it to proceed? Wait? Abort?
- Consider these cases:

$w_U(X) \; . \; . \; . \; r_T(X)$
$r_U(X) \; . \; . \; . \; w_T(X)$
$w_U(X) \; . \; . \; . \; w_T(X)$

OK

START(U), ...,START(T), ..., $w_U(X)$, ..., $r_T(X)$

Too late

START(T), ...,START(U), ..., $w_U(X)$, ..., $r_T(X)$

Timestamps

With each element X, associate

- RT(X) = the highest timestamp of any transaction U that read X
- WT(X) = the highest timestamp of any transaction U that wrote X
- C(X) = the commit bit: true when transaction with highest timestamp that wrote X committed

If element = page, then these are associated with each page X in the buffer pool

Simplified Timestamp-based Scheduling

$w_U(X)$ . . . $r_T(X)$
$r_U(X)$ . . . $w_T(X)$
$w_U(X)$ . . . $w_T(X)$

Only for transactions that do not abort

Otherwise, may result in non-recoverable schedule

Request is $r_T(X)$
?

Request is $w_T(X)$
?

Simplified Timestamp-based Scheduling

$w_U(X) \; . \; . \; . \; r_T(X)$
$r_U(X) \; . \; . \; . \; w_T(X)$
$w_U(X) \; . \; . \; . \; w_T(X)$

Only for transactions that do not abort

Otherwise, may result in non-recoverable schedule

Request is $r_T(X)$
If TS(T) < WT(X) then ROLLBACK
Else READ and update RT(X) to larger of TS(T) or RT(X)

Request is $w_T(X)$
If TS(T) < RT(X) then ROLLBACK
Else if TS(T) < WT(X) ignore write & continue (Thomas Write Rule)
Otherwise, WRITE and update WT(X) =TS(T)

Details

Read too late:

- T wants to read X, and TS(T) < WT(X)

START(T) … START(U) … $w_U(X)$ . . . $r_T(X)$

Need to rollback T !

Details

Write too late:

- T wants to write X, and TS(T) < RT(X)

START(T) … START(U) … $r_U(X)$ . . . $w_T(X)$

Need to rollback T !

Details

Write too late, but we can still handle it:

- T wants to write X, and
  TS(T) >= RT(X) but WT(X) > TS(T)

START(T) … START(V) … w_V(X) . . . w_T(X)

Don't write X at all !
(Thomas' rule)

View-Serializability

- By using Thomas' rule we do not obtain a conflict-serializable schedule

- Instead, we obtain a view-serializable schedule

Ensuring Recoverable Schedules

- Recall the definition: if a transaction reads an element, then the transaction that wrote it must have already committed
- Use the commit bit C(X) to keep track if the transaction that last wrote X has committed

Ensuring Recoverable Schedules

Read dirty data:

- T wants to read X, and WT(X) < TS(T)
- Seems OK, but…

START(U) … START(T) … w_U(X). . . r_T(X)… ABORT(U)

If C(X)=false, T needs to wait for it to become true

Ensuring Recoverable Schedules

Thomas' rule needs to be revised:

- T wants to write X, and WT(X) > TS(T)
- Seems OK not to write at all, but …

START(T) … START(U)… w_U(X). . . w_T(X)… ABORT(U)

If C(X)=false, T needs to wait for it to become true

Timestamp-based Scheduling

Request is $r_T(X)$
If TS(T) < WT(X) then ROLLBACK
Else If C(X) = false, then WAIT
Else READ and update RT(X) to larger of TS(T) or RT(X)

Request is $w_T(X)$
If TS(T) < RT(X) then ROLLBACK
Else if TS(T) < WT(X)
Then If C(X) = false then WAIT
else IGNORE write (Thomas Write Rule)
Otherwise, WRITE, and update WT(X)=TS(T), C(X)=false

Summary of Timestamp-based Scheduling

- Conflict-serializable
- Recoverable
  - Even avoids cascading aborts
- Does NOT handle phantoms

Multiversion Timestamp

- When transaction T requests r(X) but WT(X) > TS(T), then T must rollback
- Idea: keep multiple versions of X: $X_t, X_{t-1}, X_{t-2}, \ldots$

$$\boxed{TS(X_t) > TS(X_{t-1}) > TS(X_{t-2}) > \ldots}$$

- Let T read an older version, with appropriate timestamp

Details

- When $w_T(X)$ occurs,
  create a new version, denoted X_t where t = TS(T)

- When $r_T(X)$ occurs,
  find most recent version X_t such that t < TS(T)
  Notes:
  - $WT(X_t)$ = t and it never changes
  - $RT(X_t)$ must still be maintained to check legality of writes

- Can delete X_t if we have a later version X_{t1} and all active transactions T have TS(T) > t1

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{18}

$R_6(X)$ -- what happens?

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens?

$W_5(X)$ – what happens?

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{18}

$R_6(X)$ -- what happens? Return X_3

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens?

$W_5(X)$ – what happens?

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$W_5(X)$ – what happens?

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{14} X_{18}

$R_6(X)$ -- what happens? Return X_3

$W_{14}(X)$ – what happens?

$R_{15}(X)$ – what happens? Return X_{14}

$W_5(X)$ – what happens? ABORT

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{14} X_{18}

$R_6(X)$ -- what happens? Return X_3
$W_{14}(X)$ – what happens?
$R_{15}(X)$ – what happens? Return X_{14}
$W_5(X)$ – what happens? ABORT

When can we delete X_3?

Example (in class)

TS(T)=6

X_3 X_9 X_{12} X_{14} X_{18}

$R_6(X)$ -- what happens? Return X_3
$W_{14}(X)$ – what happens?
$R_{15}(X)$ – what happens? Return X_{14}
$W_5(X)$ – what happens? ABORT

When can we delete X_3? When max TS(T)≥ 9

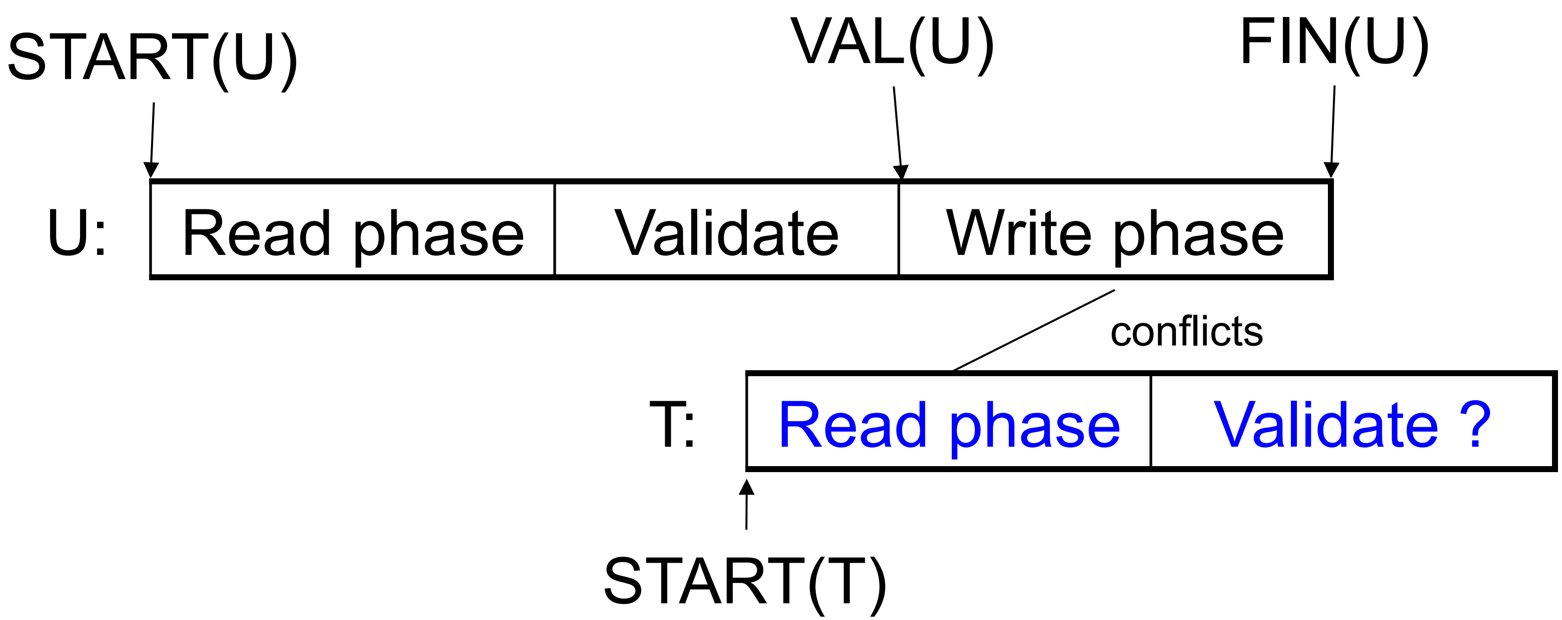
Concurrency Control by Validation

Even more optimistic than timestamp validation

- Each transaction T defines a *read set* RS(T) and a *write set* WS(T)
- Each transaction proceeds in three phases:
  - Read all elements in RS(T). Time = START(T)
  - Validate (may need to rollback). Time = VAL(T)
  - Write all elements in WS(T). Time = FIN(T)

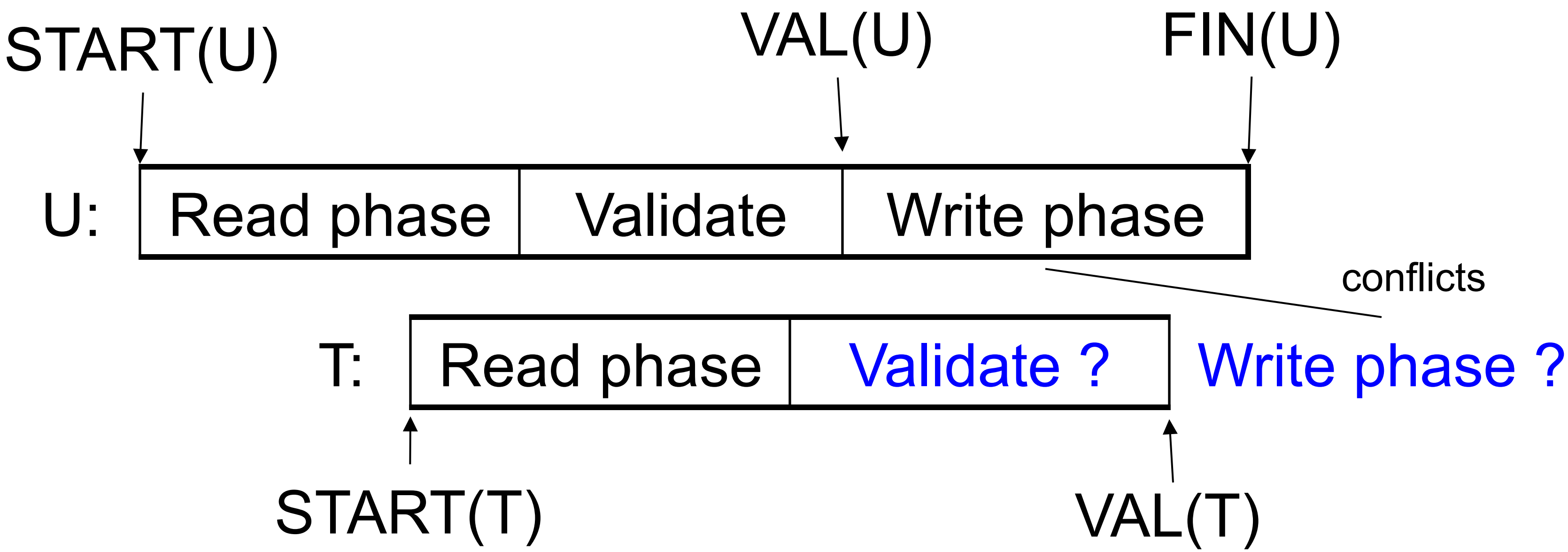
Main invariant: the serialization order is VAL(T)

Avoid $w_U(X)$ - $r_T(X)$ Conflicts



IF RS(T) ∩ WS(U) and FIN(U) > START(T)
(U has validated and U has not finished before T begun)
Then ROLLBACK(T)

Avoid $w_U(X)$ - $w_T(X)$ Conflicts



IF WS(T) ∩ WS(U) and FIN(U) > VAL(T)

(U has validated and U has not finished before T validates)

Then ROLLBACK(T)

Snapshot Isolation (SI)

- A variant of multiversion/validation
- Very efficient, and very popular
  - Oracle, PostgreSQL, SQL Server 2005
- Warning: not serializable
  - Earlier versions of postgres implemented SI for the SERIALIZABLE isolation level
  - Extension of SI to serializable has been implemented recently
  - Will discuss only the standard SI (non-serializable)

Snapshot Isolation Rules

- Each transactions receives a timestamp TS(T)
- Transaction T sees snapshot at time TS(T) of the database
- When T commits, updated pages are written to disk
- Write/write conflicts resolved by “first committer wins” rule
  - Loser gets aborted
- Read/write conflicts are ignored

Snapshot Isolation (Details)

- Multiversion concurrency control:
  - Versions of X: X_{t1}, X_{t2}, X_{t3}, . . .
- When T reads X, return $X_{TS(T)}$.
- When T writes X: if other transaction updated X, abort
  - Not faithful to “first committer” rule, because the other transaction U might have committed after T. But once we abort T, U becomes the first committer ☺

What Works and What Not

- No dirty reads (Why ?)
- No inconsistent reads (Why ?)
  - A: Each transaction reads a consistent snapshot

- No lost updates ("first committer wins")

- Moreover: no reads are ever delayed

- However: read-write conflicts not caught ! "Write skew"

Write Skew

Invariant: $X + Y \geq 0$

```
T1:
  READ(X);
  if X >= 50
      then Y = -50; WRITE(Y)
  COMMIT
```

```
T2:
  READ(Y);
  if Y >= 50
      then X = -50; WRITE(X)
  COMMIT
```

In our notation:

$R_1(X), R_2(Y), W_1(Y), W_2(X), C_1, C_2$

Starting with X=50,Y=50, we end with X=-50, Y=-50.
Non-serializable !!!

Discussions

- Snapshot isolation (SI) is like repeatable reads but also avoids some (not all) phantoms
- If DBMS runs SI and the app needs serializable:
  – use dummy writes for all reads to create write-write conflicts… but that is confusing for developers
- Recent extension of SI to make it serializable was implemented in postgres