Architecture and Implementation of Database Systems (Summer 2018)

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Part VIII

Concurrency Control

The “Hello World” of Transaction Management

- My bank issued me a debit card to access my account.
- Every once in a while, I’d use it at an ATM to draw some money from my account, causing the ATM to perform a **transaction** in the bank’s database.

```
$bal \leftarrow$ read_bal ($acct\_no$) ;
$bal \leftarrow bal - 100\,\text{CHF}$ ;
write_bal ($acct\_no$, $bal$) ;
```

- My account is properly updated to reflect the new balance.

Concurrent Access

The problem is: My wife has a card for the account, too.

- We might end up using our cards at different ATMs at the **same time**.

me	**my wife**	**DB state**
$bal \leftarrow \texttt{read}\,(acct)\,;$		1200
	$bal \leftarrow \texttt{read}\,(acct)\,;$	1200
$bal \leftarrow bal - 100\,;$		1200
	$bal \leftarrow bal - 200\,;$	1200
$\texttt{write}\,(acct\,,\,bal)\,;$		1100
	$\texttt{write}\,(acct\,,\,bal)\,;$	1000

- The first update was **lost** during this execution. Lucky me!

Another Example

- This time, I want to **transfer** money over to another account.

```
   // Subtract money from source (checking) account
1  chk_bal ← read_bal (chk_acct_no) ;
2  chk_bal ← chk_bal − 500 CHF ;
3  write_bal (chk_acct_no , chk_bal) ;
   // Credit money to the target (saving) account
4  sav_bal ← read_bal (sav_acct_no) ;
5  sav_bal ← sav_bal + 500 CHF ;
6  write_bal (sav_acct_no , sav_bal) ;
```

- Before the transaction gets to step **6**, its execution is **interrupted/cancelled** (power outage, disk failure, software bug, . . .). My money is **lost** ☹.

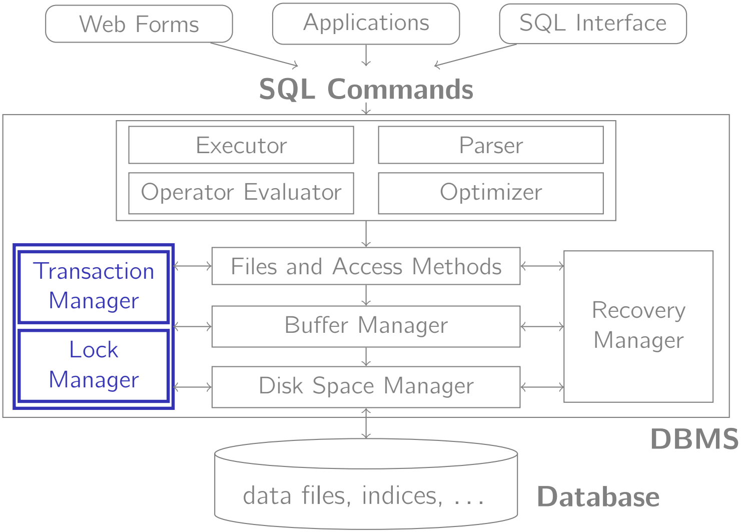
ACID Properties

One of the key benefits of a database system are the **transaction properties** guaranteed to the user:

A	Atomicity	Either **all** or **none** of the updates in a database transaction are applied.
C	Consistency	Every transaction brings the database from one **consistent** state to another.
I	Isolation	A transaction must not see any effect from other transactions that run in parallel.
D	Durability	The effects of a **successful** transaction maintain persistent and may not be undone for system reasons.

A challenge is to preserve these guarantees even with **multiple users** accessing the database **concurrently**.

Concurrency Control



Anomalies: Lost Update

- We already saw a **lost update** example on slide 296.
- The effects of one transaction are lost, because of an uncontrolled overwriting by the second transaction.

Anomalies: Inconsistent Read

Consider the money transfer example (slide 297), expressed in SQL syntax:

```
     Transaction 1                      Transaction 2
UPDATE Accounts
  SET balance = balance - 500
  WHERE customer = 4711
    AND account_type = 'C';
                                   SELECT SUM(balance)
                                     FROM Accounts
                                    WHERE customer = 4711;
UPDATE Accounts
  SET balance = balance + 500
  WHERE customer = 4711
    AND account_type = 'S';
```

- Transaction 2 sees an **inconsistent** database state.

Anomalies: Dirty Read

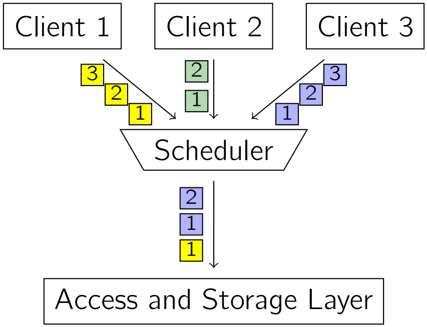
At a different day, my wife and me again end up in front of an ATM at roughly the same time:

me	**my wife**	**DB state**
$bal \leftarrow$ `read` $(acct)$;		1200
$bal \leftarrow bal - 100$;		1200
`write` $(acct, bal)$;		1100
	$bal \leftarrow$ `read` $(acct)$;	1100
	$bal \leftarrow bal - 200$;	1100
`abort` ;		1200
	`write` $(acct, bal)$;	900

- My wife's transaction has already read the modified account balance before my transaction was **rolled back**.

Concurrent Execution

- The **scheduler** decides the execution order of concurrent database accesses.



Database Objects and Accesses

- We now assume a slightly simplified model of database access:
  1. A database consists of a number of named **objects**. In a given database state, each object has a **value**.
  2. Transactions access an object o using the two operations `read` o and `write` o.
- In a **relational** DBMS we have that

$$\text{object} \equiv \text{attribute} \quad .$$

Transactions

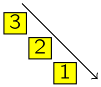
A **database transaction** T is a (strictly ordered) sequence of **steps**. Each **step** is a pair of an **access operation** applied to an **object**.

- Transaction $T = \langle s_1, \dots, s_n \rangle$
- Step $s_i = (a_i, e_i)$
- Access operation $a_i \in \{\texttt{r(ead)}, \texttt{w(rite)}\}$

The **length** of a transaction T is its number of steps $|T| = n$.

We could write the money transfer transaction as

$$T = \langle\ (\texttt{read}, \mathit{Checking}), (\texttt{write}, \mathit{Checking}),\ (\texttt{read}, \mathit{Saving}), (\texttt{write}, \mathit{Saving})\ \rangle$$



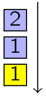
or, more concisely,

$$T = \langle r(C), w(C), r(S), w(S) \rangle \enspace .$$

Schedules

A **schedule** S for a given set of transactions $\mathbf{T} = \{T_1, \ldots, T_n\}$ is an arbitrary sequence of execution steps

$$S(k) = (T_j, a_i, e_i) \qquad k = 1 \ldots m \ ,$$



such that

1. S contains all steps of all transactions an nothing else and
2. the order among steps in each transaction T_j is preserved:

$$(a_p, e_p) < (a_q, e_q) \text{ in } T_j \Rightarrow (T_j, a_p, e_p) < (T_j, a_q, e_q) \text{ in } S \ .$$

We sometimes write

$$S = \langle r_1(B), r_2(B), w_1(B), w_2(B) \rangle$$

to mean

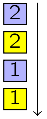
$$\begin{array}{ll} S(1) = (T_1, \texttt{read}, B) & S(3) = (T_1, \texttt{write}, B) \\ S(2) = (T_2, \texttt{read}, B) & S(4) = (T_2, \texttt{write}, B) \end{array}$$

Serial Execution

One particular schedule is **serial execution**.

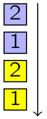
- A schedule S is **serial** iff, for each contained transaction T_j, all its steps follow each other (no interleaving of transactions).

Consider again the ATM example from slide 296.



- $S = \langle r_1(B), r_2(B), w_1(B), w_2(B) \rangle$
- This schedule is **not** serial.

If my wife had gone to the bank one hour later, "our" schedule probably would have been serial.



- $S = \langle r_1(B), w_1(B), r_2(B), w_2(B) \rangle$

Correctness of Serial Execution

- Anomalies such as the "lost update" problem on slide 296 can **only** occur in multi-user mode.
- If all transactions were fully executed one after another (no concurrency), no anomalies would occur.
- **Any serial execution is correct.**

- Disallowing concurrent access, however, is **not practical**.
- Therefore, allow concurrent executions if they are **equivalent** to a serial execution.

Conflicts

What does it mean for a schedule S to be equivalent to another schedule S'?

- Sometimes, we may be able to **reorder** steps in a schedule.
  - We must not change the order among steps of any transaction T_j ($\nearrow$ slide 306).
  - Rearranging operations must not lead to a different **result**.
- Two operations (a, e) and (a', e') are said to be **in conflict** $(a, e) \nleftrightarrow (a', e')$ if their order of execution matters.
  - When reordering a schedule, we must not change the relative order of such operations.
- Any schedule S' that can be obtained this way from S is said to be **conflict equivalent** to S.

Conflicts

Based on our `read`/`write` model, we can come up with a more machine-friendly definition of a conflict.

- Two operations (T_i, a, e) and (T_j, a', e') are **in conflict** in S if
  1. they belong to two **different transactions** $(T_i \neq T_j)$,
  2. they access the **same database object**, *i.e.*, $e = e'$, and
  3. at least one of them is a `write` operation.
- This inspires the following conflict matrix:

	`read`	`write`
`read`		×
`write`	×	×

- **Conflict relation $\prec_S$:**

$$(T_i, a, e) \prec_S (T_j, a', e')$$
$$:=$$
$$(a, e) \not\leftrightarrow (a', e') \wedge (T_i, a, e) \text{ occurs before } (T_j, a', e') \text{ in } S \wedge T_i \neq T_j$$

Conflict Serializability

- A schedule S is **conflict serializable** iff it is conflict equivalent to **some** serial schedule S'.
- **The execution of a conflict-serializable S schedule is correct.**
  - S does **not** have to be a serial schedule.
- This allows us to **prove** the correctness of a schedule S based on its **conflict graph** $G(S)$ (also: **serialization graph**).
  - **Nodes** are all transactions T_i in S.
  - There is an **edge** $T_i \rightarrow T_j$ iff S contains operations (T_i, a, e) and (T_j, a', e') such that $(T_i, a, e) \prec_S (T_j, a', e')$.
- S is conflict serializable if $G(S)$ is **acyclic**.[18]

[18] A serial execution of S could be obtained by sorting $G(S)$ **topologically**.

Serialization Graph

Example: ATM transactions ($\nearrow$ slide 296)

- $S = \langle r_1(A), r_2(A), w_1(A), w_2(A) \rangle$
- Conflict relation:
  $(T_1, \mathtt{r}, A) \prec_S (T_2, \mathtt{w}, A)$
  $(T_2, \mathtt{r}, A) \prec_S (T_1, \mathtt{w}, A)$
  $(T_1, \mathtt{w}, A) \prec_S (T_2, \mathtt{w}, A)$

$$\begin{tikzcd}
 & T_2 \arrow[dl, bend left] \\
T_1 \arrow[ur, bend left] \arrow[ur, phantom, "\circlearrowright"] &
\end{tikzcd}$$

$\rightarrow$ **not** serializable

Example: Two money transfers ($\nearrow$ slide 297)

- $S = \langle r_1(C), w_1(C), r_2(C), w_2(C), r_1(S), w_1(S), r_2(S), w_2(S) \rangle$
- Conflict relation:
  $(T_1, \mathtt{r}, C) \prec_S (T_2, \mathtt{w}, C)$
  $(T_1, \mathtt{w}, C) \prec_S (T_2, \mathtt{r}, C)$
  $(T_1, \mathtt{w}, C) \prec_S (T_2, \mathtt{w}, C)$
  $\vdots$

$$\begin{tikzcd}
 & T_2 \\
T_1 \arrow[ur, bend left] &
\end{tikzcd}$$

$\rightarrow$ serializable

Query Scheduling

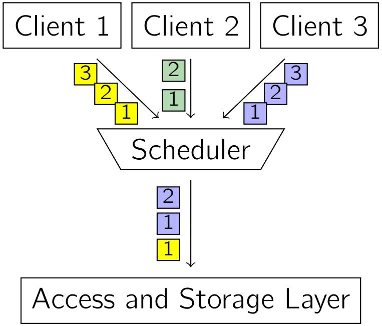
Can we build a scheduler that **always** emits a serializable schedule?

Idea:

- Require each transaction to obtain a **lock** before it accesses a data object o:

```
lock o ;
...access o ...;
unlock o ;
```

- This prevents **concurrent** access to o.



Locking

- If a lock cannot be granted (*e.g.*, because another transaction T' already holds a **conflicting** lock) the requesting transaction T_i gets **blocked**.
- The scheduler **suspends** execution of the blocked transaction T.
- Once T' **releases** its lock, it may be granted to T, whose execution is then **resumed**.
- Since other transactions can continue execution while T is blocked, locks can be used to **control the relative order of operations**.

Locking and Serializability

✎ **Does locking guarantee serializable schedules, yet?**

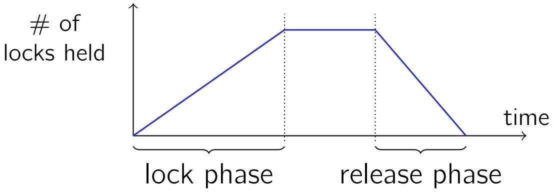
ATM Transaction with Locking

Transaction 1	**Transaction 2**	**DB state**
`lock` (*acct*) ;		1200
`read` (*acct*) ;		
`unlock` (*acct*) ;		
	`lock` (*acct*) ;	
	`read` (*acct*) ;	
	`unlock` (*acct*) ;	
`lock` (*acct*) ;		
`write` (*acct*) ;		1100
`unlock` (*acct*) ;		
	`lock` (*acct*) ;	
	`write` (*acct*) ;	1000
	`unlock` (*acct*) ;	

Two-Phase Locking (2PL)

The **two-phase locking protocol** poses an additional restriction:

- Once a transaction has **released** any lock, it must **not** acquire any new lock.



- Two-phase locking is **the** concurrency control protocol used in database systems today.

Again: ATM Transaction

Transaction 1	**Transaction 2**	**DB state**
`lock` (*acct*) ;		1200
`read` (*acct*) ;		
`unlock` (*acct*) ;		
	`lock` (*acct*) ;	
	`read` (*acct*) ;	
	`unlock` (*acct*) ;	
`lock` (*acct*) ; ↯		
`write` (*acct*) ;		1100
`unlock` (*acct*) ;		
	`lock` (*acct*) ; ↯	
	`write` (*acct*) ;	1000
	`unlock` (*acct*) ;	

A 2PL-Compliant ATM Transaction

- To comply with the two-phase locking protocol, the ATM transaction must not acquire any new locks after a first lock has been released.

```
lock (acct) ;                      } lock phase
bal ← read_bal (acct) ;
bal ← bal − 100 CHF ;
write_bal (acct, bal) ;
unlock (acct) ;                    } unlock phase
```

Resulting Schedule

Transaction 1	Transaction 2	DB state
`lock` (*acct*) ;		1200
`read` (*acct*) ;		
	`lock` (*acct*) ;	
`write` (*acct*) ;	↓ Transaction	1100
`unlock` (*acct*) ;	blocked	
	`read` (*acct*) ;	
	`write` (*acct*) ;	900
	`unlock` (*acct*) ;	

- The use of locking lead to a correct (and serializable) schedule.

Lock Modes

- We saw earlier that two **read** operations do not conflict with each other.
- Systems typically use different types of locks (**"lock modes"**) to allow read operations to run concurrently.
  - **read locks** or **shared locks**: mode `S`
  - **write locks** or **exclusive locks**: mode `X`
- Locks are only in conflict if at least one of them is an `X` lock:

	shared (`S`)	exclusive (`X`)
shared (`S`)		×
exclusive (`X`)	×	×

- It is a safe operation in two-phase locking to **convert** a shared lock into an exclusive lock during the lock phase.

Deadlocks

- Like many lock-based protocols, two-phase locking has the risk of **deadlock** situations:

Transaction 1	**Transaction 2**
`lock` (A) ;	
⋮	`lock` (B)
do something	⋮
⋮	do something
`lock` (B) [wait for T_2 to release lock]	⋮
	`lock` (A) [wait for T_1 to release lock]

- Both transactions would wait for each other **indefinitely**.

Deadlock Handling

A typical approach to deal with deadlocks is **deadlock detection**:

- The system maintains a **waits-for graph**, where an edge $T_1 \rightarrow T_2$ indicates that T_1 is blocked by a lock held by T_2.
- Periodically, the system tests for **cycles** in the graph.
- If a cycle is detected, the deadlock is **resolved** by **aborting** one or more transactions.
- Selecting the **victim** is a challenge:
  - Blocking **young** transactions may lead to **starvation**: the same transaction is cancelled again and again.
  - Blocking an **old** transaction may cause a lot of investment to be thrown away.

Deadlock Handling

Other common techniques:

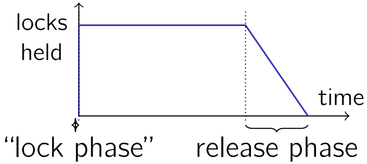
- **Deadlock prevention:** *e.g.*, by treating handling lock requests in an **asymmetric** way:
  - **wait-die**: A transaction is never blocked by an **older** transaction.
  - **wound-wait**: A transaction is never blocked by a **younger** transaction.
- **Timeout:** Only wait for a lock until a timeout expires. Otherwise assume that a deadlock has occurred and **abort**.

E.g., IBM DB2 UDB:

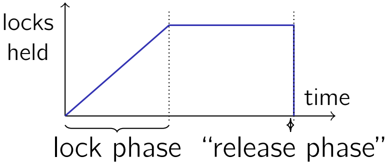
```
db2 => GET DATABASE CONFIGURATION;
  :
  :
Interval for checking deadlock (ms)       (DLCHKTIME) = 10000
Lock timeout (sec)                      (LOCKTIMEOUT) = -1
```

Variants of Two-Phase Locking

- The two-phase locking protocol does not prescribe exactly when locks have to acquired and released.
- Possible variants:



preclaiming 2PL

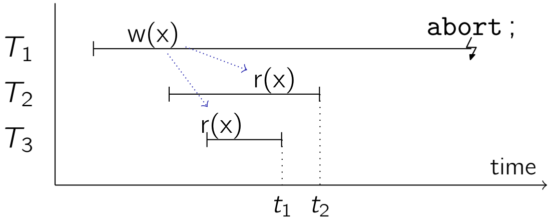


strict 2PL

- ✎ **What could motivate either variant?**

Cascading Rollbacks

Consider three transactions:



- When transaction T_1 aborts, transactions T_2 and T_3 have already read data written by T_1 ($\nearrow$ dirty read, slide 302)
- T_2 and T_3 need to be **rolled back**, too.
- T_2 and T_3 **cannot** commit until the fate of T_1 is known.
- two-phase locking vs. strict two-phase locking

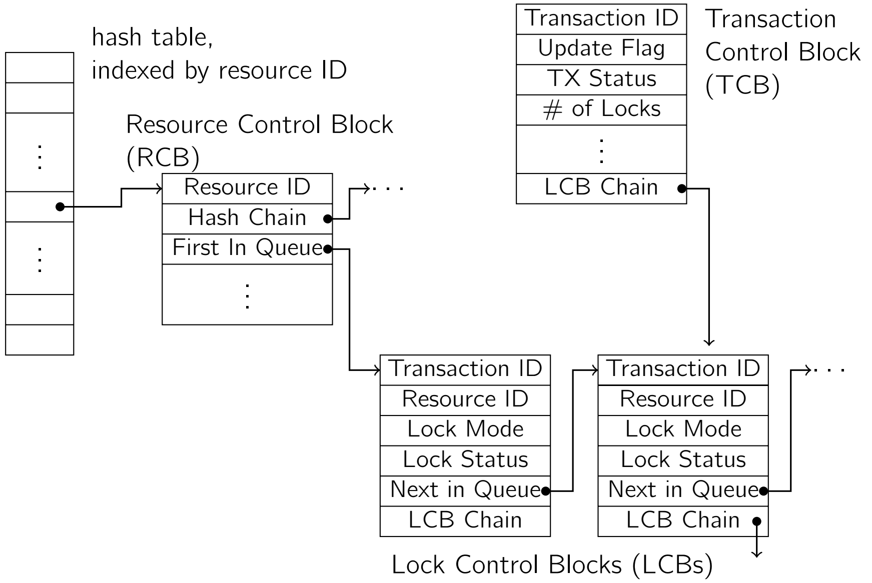
Implementing a Lock Manager

We'd like the Lock Manager to do three tasks very efficiently:

1. Check which locks are currently held for a given **resource** (in order to decide whether another lock request can be granted).
2. When a lock is released, **transactions** that **requested** locks on the **same resource** have to be identified and granted the lock.
3. When a transaction **terminates**, all held locks must be released.

What is a good data structure to accommodate these needs?

Bookkeeping

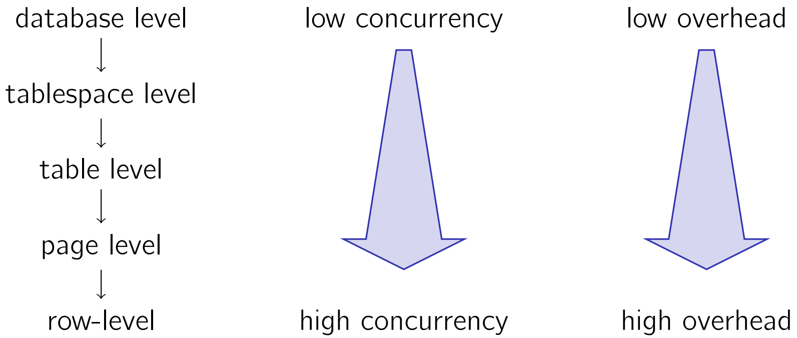


Implementing Lock Manager Tasks

1. The locks held for a given **resource** can be found using a **hash lookup**.
   - Linked list of Lock Control Blocks via ‘First In Queue’/‘Next in Queue’
   - The list contains **all** lock requests, granted or not.
   - The transaction(s) at the **head** of the list are the ones that currently hold a lock on the resource.
2. When a lock is **released** (*i.e.*, its LCB removed from the list), the next transaction(s) in the list are considered for granting the lock.
3. All locks held by a single **transaction** can be identified via the linked list ‘LCB Chain’ (and easily released upon transaction termination).

Granularity of Locking

The **granularity** of locking is a trade-off:



- **Idea: multi-granularity** locking

Multi-Granularity Locking

- Decide the granularity of locks held **for each transaction** (depending on the characteristics of the transaction).
  - A row lock, *e.g.*, for

```
SELECT * FROM CUSTOMERS                                   Q1
 WHERE C_CUSTKEY = 42
```

    and a table lock for

```
SELECT * FROM CUSTOMERS                                   Q2
```

- How do such transactions know about each others' locks?
  - Note that locking is **performance-critical**. Q_2 doesn't want to do an extensive search for row-level conflicts.

Intention Locks

Databases use an additional type of locks: **intention locks**.

- Lock mode **intention share**: `IS`
- Lock mode **intention exclusive**: `IX`
- Conflict matrix:

	S	X	IS	IX
S		×		×
X	×	×	×	×
IS		×		
IX	×	×		

- A lock `I`□ on a coarser level means that there's some □ lock on a lower level.

Intention Locks

Protocol for multi-granularity locking:

1. A transaction can lock any granule g in $\Box \in \{\texttt{S}, \texttt{X}\}$ mode.
2. Before a granule g can be locked in $\Box$ mode, it has to obtain an $\texttt{I}\Box$ lock on **all** coarser granularities than contain g.

Query Q_1 would, *e.g.*,

- obtain an `IS` lock on **table** `CUSTOMERS`
  (also on on tablespace and database) and
- obtain an `S` lock on the **tuple(s)** with `C_CUSTKEY = 42`.

Query Q_2 would place an

- `S` lock on `table CUSTOMERS`
  (and an `IS` lock on tablespace and database).

Detecting Conflicts

Now suppose a write query comes in:

```
UPDATE CUSTOMERS                                    Q3
  SET NAME = 'John Doe'
  WHERE C_CUSTKEY = 17
```

It'll want to place

- an `IX` lock on **table** `CUSTOMER` (and . . .) and
- an `X` lock on the **row** holding customer `17`.

As such it is

- **compatible** with Q_1
  (there's no conflict between `IX` and `IS` on the table level),
- but **incompatible** with Q_2
  (the `S` lock held by Q_2 is in **conflict** with Q_3's `IX` lock).

Consistency Guarantees and SQL 92

Sometimes, some degree of inconsistency may be acceptable for specific applications:

- "Mistakes" in few data sets, *e.g.*, will not considerably affect the outcome of an aggregate over a huge table.
  - $\rightsquigarrow$ Inconsistent read anomaly
- SQL 92 specifies different **isolation levels**.
- *E.g.*,

```
SET ISOLATION SERIALIZABLE;
```

- Obviously, less strict consistency guarantees should lead to increased throughput.

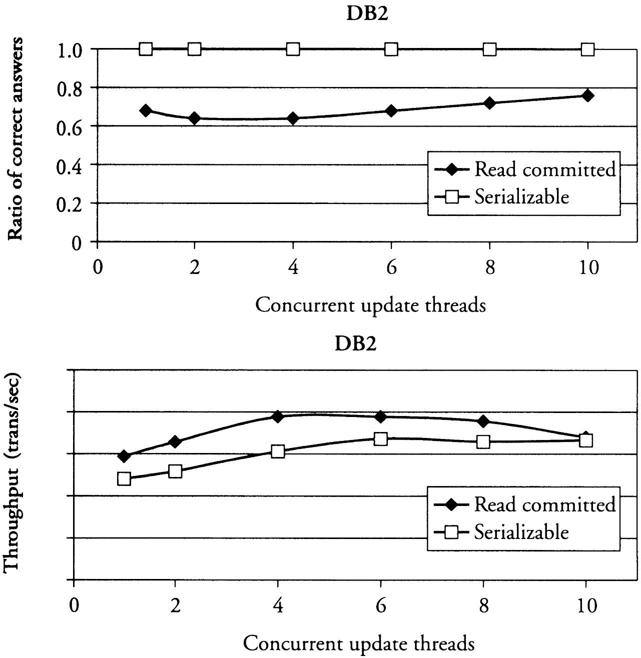
SQL 92 Isolation Levels

read uncommitted (also: ‘dirty read’ or ‘browse’)
: Only **write locks** are acquired (according to strict 2PL).

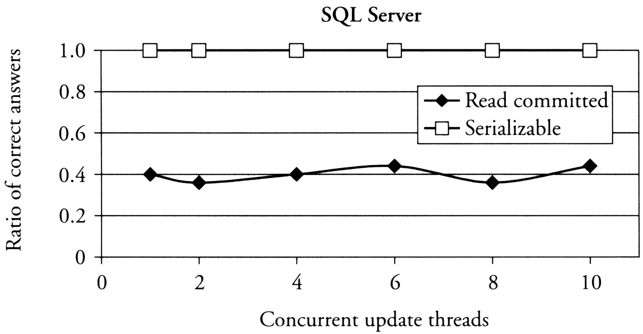
read committed (also: ‘cursor stability’)
: **Read locks** are only held for as long as a cursor sits on the particular row. **Write locks** acquired according to strict 2PL.

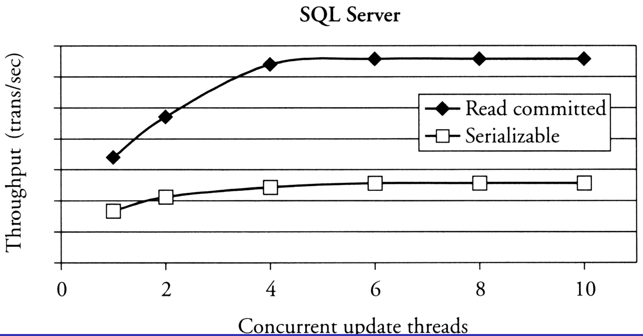
repeatable read (also: ‘read stability’)
: Acquires **read** and **write locks** according to strict 2PL.

serializable
: Additionally obtains locks to avoid **phantom reads**.



Dennis Shasha nad Philippe Bonnet. Database Tuning. Morgan Kaufmann, 2003.





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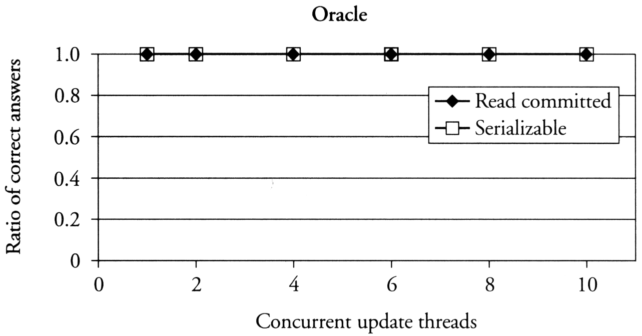
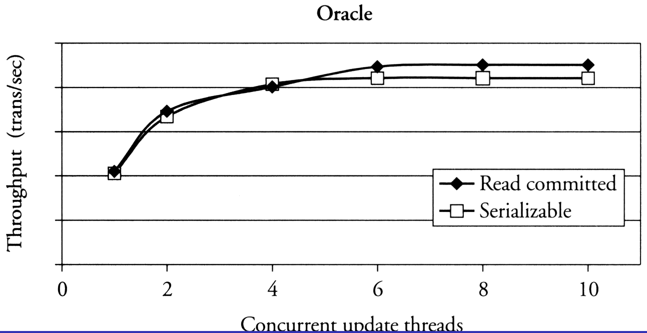




Dennis Shasha nad Philippe Bonnet. Database Tuning. Morgan Kaufmann, 2003.

Resulting Consistency Guarantees

isolation level	**dirty read**	**non-repeat. rd**	**phantom rd**
read uncommitted	possible	possible	possible
read committed	not possible	possible	possible
repeatable read	not possible	not possible	possible
serializable	not possible	not possible	not possible

- Some implementations support more, less, or different levels of isolation.
- Few applications really need serializability.

Locking and B-Trees

Transaction 1	**Transaction 2**	**Result**

`SELECT COUNT (*)`
`FROM Customers`
`WHERE Name = 'Sam'`
		2

	`INSERT INTO Customers`
	`VALUES (..., 'Sam', ...)`
		ok

`SELECT COUNT (*)`
`FROM Customers`
`WHERE Name = 'Sam'`
		3 $\lightning$

Transaction 1 "sees" the concurrent insert done by Transaction 2.

$\rightarrow$ **Isolation property violated.**

This is an instance of the **phantom problem**.

Avoiding Phantoms

Locking only tuples cannot avoid the phantom problem.

- The tuple added by T_2 is new; T_1 could never have locked it before.
- To avoid the phantom problem, we also have to lock **absent** tuples.

Phantoms can be avoided with:

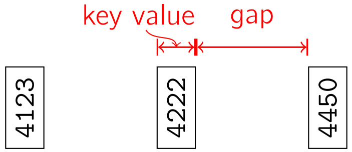
- **Predicate Locking:** For each query, lock the predicates that it uses.

  Representing, finding, and comparing predicates can be difficult and inefficient.

- **Key-Range Locking:** Lock **index entries** that match the predicate.
  - *E.g.*, in the previous example, lock the index key `Sam`.

Key-Range Locking

- Use B-trees to lock **key values**, **not tuples**!
  - → This is somewhat orthogonal to regular data locking.
- In general, we want to lock **ranges** of key values.
  - → Including **absence** of key values.
  - → Lock existing **key values** and **gaps**.

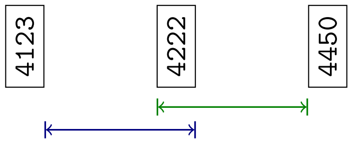


→ The current index content determines which ranges can be locked.

Key-Range Locking

Typically:

- Acquire **one lock** to mean a key value **and** its neighboring gap:



→ **Previous key locking:** $\vdash\!\leftrightarrow\!\dashv$
Lock covers key value *x* and the gap that **follows** *x*.

→ **Next key locking:** $\vdash\!\leftrightarrow\!\dashv$
Lock covers key value *x* and the gap that **precedes** *x*.

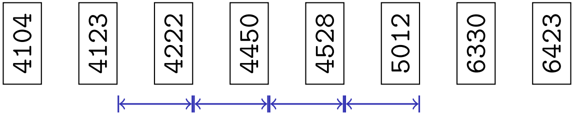
This way, existing key values can be used as lookup keys in the system's **lock manager** (which is typically organized as a hash table).

Reading Transactions

Idea:

- Queries acquire `S` locks for all key ranges that **intersect** with ranges in query predicates.

E.g., scan range [4200, 5000]:

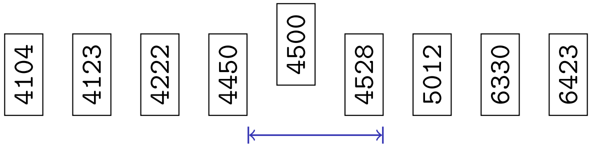


→ Ranges]4123, 4200[and]5000, 5012] locked "too much"!

Inserts

- Inserts need to acquire a lock on the **gap** into which they want to insert.
- Thus, with next key locking: acquire lock on next-largest key.

E.g., insert 4500:



→ Acquire `X` lock on 4528 (which covers range $]4450, 4528]$).

→ If the reading transaction from the previous slide still holds its locks, a conflict on 4528 will be detected (and the insert will have to wait).

→ Insert new key and `X` lock it immediately.

Lock Duration

Readers:

- Keep the range locked until the transaction commits.
  This is to make sure the range can be re-read at any time without seeing phantoms.

Inserts:

- Keep **newly inserted entry** `X` locked until commit time.
  → This prevents others from reading un-committed data.
- The lock on the **next key** (4528 here), however, can be released **immediately**.
  → Acquiring the lock with "instant duration" ensures there is no co-running reader for that range.
  → Once the new key is inserted, readers (or writers) are free to lock the next key (4528), since its associated range (]4500, 4528] now) only covers the gap **without** the newly inserted key.

Why Bother?

This ability to lock with instant duration is very relevant in practice.

- Inserts at the **right end** of a B-tree are a very common pattern.
  - → Next key locking requires an extra $+\infty$ index entry, by the way.
  - → All append queries will lock this $+\infty$ entry.
  - → When the lock on $+\infty$ is an instant lock, other inserts can proceed immediately.

→ Note how this also favors next key locking over previous key locking.

Deletions

To **delete** an entry *x*, the transaction has to obtain

- an `X` lock on the to-be-deleted entry *x*,
  - → Make sure no other transaction still depends on *x*.
  - → The lock is effectively instant, since the transaction is about to remove *x* anyway.
- an `X` lock on *x*'s **next key** until **commit time**.
  - → ✎ **Why?**

Key-Value Locks in Practice

IBM DB2 does not lock index entries explicitly.

- Instead, DB2 performs **data-only locking**.
- A locked tuple **implies** a key-range lock in **all** indexes on the table.
- When checking for lock compatibility, DB2 looks for already held locks, but also considers the **isolation level** of the lock holder.

Data-only locking may lead to unexpected **side effects**:

- *E.g.*, a scan criterion on one column may lead to locks in scattered regions of other attributes.

On the positive side, deriving key-range locks from row locks reduces the number of locks to maintain (and thus the complexity of the lock manager).

Ghost Records

Support for **ghost records** may ease key-range locking considerably.

- **Deletes** will not actually remove the index entry, but only turn the record into a ghost.
- The ghost still represents a valid range boundary (locks can be acquired on ghosts just as on normal records).
- Flipping the ghost bit is merely a form of **value update** of the record.
  - → Value updates do not need range locks as long as they do not modify the key value.

Ghost Records and Inserts

The same advantages also hold for **inserts** if a ghost with the right key value already exists.

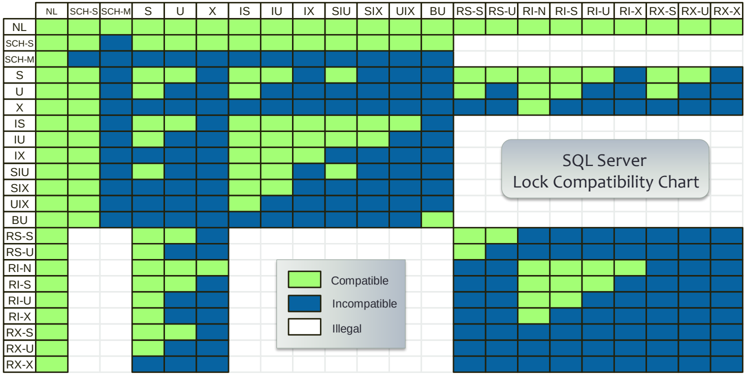
→ Need to lock only the key value itself (neighboring range is often implicit, but not strictly required).

Existence of a matching ghost need **not** be a coincidence.

Trick:

- Invoke a short, **separate transaction** that creates the ghost for us.
- The transaction will have to acquire range locks. But it will commit immediately (and release its locks).

Locking in Practice—SQL Server



SQL Server Lock Compatibility Chart

	NL	SCH-S	SCH-M	S	U	X	IS	IU	IX	SIU	SIX	UIX	BU	RS-S	RS-U	RI-N	RI-S	RI-U	RI-X	RX-S	RX-U	RX-X
NL	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩
SCH-S	🟩	🟩	🟦	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	🟩	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
SCH-M	🟩	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
S	🟩	🟩	🟦	🟩	🟩	🟦	🟩	🟩	🟦	🟩	🟦	🟦	🟦	🟩	🟩	🟩	🟩	🟩	🟦	🟩	🟩	🟦
U	🟩	🟩	🟦	🟩	🟦	🟦	🟩	🟦	🟦	🟦	🟦	🟦	🟦	🟩	🟦	🟩	🟩	🟦	🟦	🟩	🟦	🟦
X	🟩	🟩	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟩	🟦	🟦	🟦	🟦	🟦	🟦
IS	🟩	🟩	🟦	🟩	🟩	🟦	🟩	🟩	🟩	🟩	🟩	🟩	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
IU	🟩	🟩	🟦	🟩	🟦	🟦	🟩	🟩	🟩	🟩	🟩	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
IX	🟩	🟩	🟦	🟦	🟦	🟦	🟩	🟩	🟩	🟦	🟦	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
SIU	🟩	🟩	🟦	🟩	🟦	🟦	🟩	🟩	🟦	🟩	🟦	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
SIX	🟩	🟩	🟦	🟦	🟦	🟦	🟩	🟩	🟦	🟦	🟦	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
UIX	🟩	🟩	🟦	🟦	🟦	🟦	🟩	🟦	🟦	🟦	🟦	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
BU	🟩	🟩	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟩	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜	⬜
RS-S	🟩	⬜	⬜	🟩	🟩	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟩	🟩	🟦	🟦	🟦	🟦	🟦	🟦	🟦
RS-U	🟩	⬜	⬜	🟩	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟩	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦
RI-N	🟩	⬜	⬜	🟩	🟩	🟩	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟩	🟩	🟩	🟩	🟦	🟦	🟦
RI-S	🟩	⬜	⬜	🟩	🟩	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟩	🟩	🟩	🟦	🟦	🟦	🟦
RI-U	🟩	⬜	⬜	🟩	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟩	🟩	🟦	🟦	🟦	🟦	🟦
RI-X	🟩	⬜	⬜	🟩	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟩	🟦	🟦	🟦	🟦	🟦	🟦
RX-S	🟩	⬜	⬜	🟩	🟩	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦
RX-U	🟩	⬜	⬜	🟩	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦
RX-X	🟩	⬜	⬜	🟦	🟦	🟦	⬜	⬜	⬜	⬜	⬜	⬜	⬜	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦	🟦

🟩	Compatible
🟦	Incompatible
⬜	Illegal

NL	No Lock	IX	Intent Exclusive	RI-S	Insert Range-Shared
SCH-S	Schema Stability Lock	SIU	Share with Intent Update	RI-U	Insert Range-Update
SCH-M	Schema Modification Lock	SIX	Share with Intent Exclusive	RI-X	Insert Range-Exclusive
S	Shared	UIX	Update with Intent Exclusive	RX-S	Exclusive Range-Shared
U	Update	BU	Bulk Update	RX-U	Exclusive Range-Update
X	Exclusive	RS-S	Shared Range-Shared	RX-X	Exclusive Range-Exclusive
IS	Intent Shared	RS-U	Shared Range-Update		
IU	Intent Update	RI-N	Insert Range-Null		

sqlbits READPAST & Furious: Locking Blocking and Isolation · Mark Broadbent · sqlcloud.co.uk

Multi-User and Multi-Thread Support

So far we looked at ill effects **between user transactions**.

$\rightarrow$ **Locks on data objects** helped to **isolate transactions**.

Parallel threads might cause additional problems:

$\rightarrow$ Two writers, different data objects, same page $\leadsto$ **corrupted data**.

$\rightarrow$ Locks will not isolate threads that belong to the **same transaction**.

$\rightarrow$ How do we protect **internal data structures** (lock table, buffer pool, etc.)?

- Lock manager can only lock user data objects!

This calls for a mechanism to **isolate threads** (not transactions).

$\rightarrow$ Short-lived, in-memory "locks" or **latches**.
(The term "lock" is reserved for transaction-level locking.)

Page Latches

Latches protect data at a **page granularity**.

→ This has also been called **storage-layer concurrency**.

To achieve **high concurrency**:

- Hold latches **as short as possible**.
- Hold **few latches only** (and/or latch at fine granularities).

In addition:

- Choose a **fast implementation** for latches.
  - → no frills like **deadlock checking**
  - → instead: avoid deadlocks by **coding discipline**

Page Latches (Data Pages)

Example:

- Latches on data pages make page modifications appear as an **atomic operation**.
  - → Protect from, *e.g.*, observing a corrupt page.
- Latching is **in-memory only**.
  - → No I/O while holding a page latch.
  - → Latches are not flushed to disk.
- Only hold **one latch at a time**.
  - → ✎ **Why?**

Latches and B-Trees (Search)

✎ **Order for latch acquisition/release during a B-tree search?**

Latch Coupling

With latch coupling, a thread may hold more than one latch at a time.

→ A **deadlock** still cannot occur:

- Every thread will navigate/acquire latches top-down.
- All threads acquire latches in same order → no deadlock.

Latches and B-Trees (Updates)

Updates to B-trees operate **bottom-up**.

Possible strategy:

- Acquire read latches as during search, but **keep** all latches.
  - → Ensure that the parent (grandparent, ...) is still the parent during bottom-up processing.
- Acquire **write latches bottom-up**.
  - → Latch conversion: read latch $\rightsquigarrow$ write latch.
  - → Write-latch parent before splitting a child.
- Release write latches when all necessary changes to the page are applied; release ancestor read latches when no more splits are necessary.

If the B-tree implementation uses **sibling pointers**, additional locks may have to be acquired on **sibling nodes**.

Latches and B-Trees (Updates)

The strategy on the previous slide guarantees **correctness**.

- All tree modifications are write-latched, and released latches always leave behind a consistent B-tree.

But:

The strategy entails a danger of **deadlocks**.

- **Searches** acquire their latches **top-down**.
- **Updates** acquire their (write) latches **bottom-up**.

Remember: We want latches to be lightweight $\rightarrow$ no deadlock checking.

Latching B-Trees Without Deadlocks

Deadlocks can be **avoided** when **all** operations acquire latches either top-down or bottom-up.

Thus:

- Let insert operations acquire write latches right away.

✎ **What do you think of this strategy?**

Latching B-Trees Without Deadlocks

Chances that a write latch on a parent is actually needed are **really low**.

$\rightarrow$ *E.g.*, B-tree with up to 100 entries/node $\rightarrow$ chance of a split: 2 %

Idea: (Try to) keep write latch only when really necessary.

- During tree descent, observe **space utilization** in visited nodes.
- When a node *n* has **enough space** to hold another entry, *n* definitely won't have to be split.
- For such nodes *n*, the **parent node** *p* will not have to be updated.
  - $\rightarrow$ *p* is then called **split safe**.
- The latch on that parent *p* can be released safely.

Lock Coupling Protocol (Variant 1)

readers

```
place S lock on root ;
current ← root ;
while current is not a leaf node do
    place S lock on appropriate son of current ;
    release S lock on current ;
    current ← son of current ;
```

writers

```
place X lock on root ;
current ← root ;
while current is not a leaf node do
    place X lock on appropriate son of current ;
    current ← son of current ;
    if current is safe then
        release all locks held on ancestors of current ;
```

Increasing Concurrency for Common Scenarios

- Even with lock coupling there's a considerable amount of locks on inner tree nodes (reducing concurrency).
- Chances that inner nodes are actually affected by updates are **very small**.
  - Back-of-the-envelope calculation:
    $d = 50 \Rightarrow$ every 50th insert causes a split ($2\,\%$ chance).
- An insert transaction could thus optimistically assume that no leaf split is going to happen.
  - On inner nodes, only read locks acquired during tree navigation (plus a write lock on the affected leaf).
  - If assumption is wrong, re-traverse the tree and obtain write locks.

Lock Coupling Protocol (Variant 2)

Modified protocol for **writers**:[19]

```
place S lock on root ;
current ← root ;
while current is not a leaf node do
    son ← appropriate son of current ;
    if son is a leaf then
        place X lock on son ;
    else
        place S lock on son ;
    release lock on current ;
    current ← son ;
if current is unsafe then
    release all locks and repeat with protocol Variant 1 ;
```

[19]Reader protocol remains unchanged.

B-Tree Latching and High Concurrency

- Deciding **split safety** can be difficult for **variable-length keys**.
- The strategy on the previous slide thus has to be **very conservative**.
- Effectively, many latches are still held **unnecessarily**.

Ways to improve concurrency (by holding fewer latches):

- **split proactively**: When a node is not split safe, split it right away. At least the system then suffers the unnecessary latch only once.
- **repeated root-to-leaf passes**: Descend with only read latches first. Re-traverse the tree with full write latches when a split is necessary.
- **giveup technique**: hold only single-node read latches (and risk inconsistencies); detect conflicts and re-traverse in case of a conflict.
- **B$^{\text{link}}$-trees**: slightly relax some B-tree rules.

Giveup Technique

A deadlock can only arise when a thread acquires (or tries to) a new latch before releasing an old one.

→ A thread that always only holds a **single latch** at a time can never deadlock.

Search routine with only a single latch held at any time:

```
$n \leftarrow$ root page ;
while $n$ is not a leaf do
    read-latch $n$ ;
    determine child $n'$ of $n$ ;
    un-latch $n$ ;
    $n \leftarrow n'$ ;
read-latch $n$ ;
return matching records (if any)
  ;
un-latch $n$ ;
```

Giveup Technique

There is a **risk of inconsistencies** when only a single latch is held.

- Between determining the child page n' and latching it, a **concurrent update** might have split n'.
- The search might **miss** an entry that is now on a new page.

Thus: Detect when a conflicting update has happened.

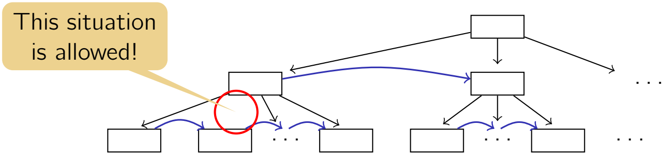
- When descending, remember the two **separator keys** k_{min} and k_{max} in n that guided to n'.
- When looking at n', first check whether k_{min} and k_{max} are still the correct separator keys for that page.
  - $\rightarrow$ Keep copies of parent's separator keys in each node.
  - $\rightarrow$ Such copies are also called **fence keys**.
- If a conflict is detected, **abort and re-try** a moment later.

B$^{\text{link}}$-Trees

Lehman and Yao[20] proposed a B-tree variant, usually referred to as **B$^{\text{link}}$-tree**, where writes must latch at most two nodes at a time.

Idea:

- Assume a B-tree with **forward sibling pointers**.
- **Relax B-tree structure**: Allow parent $\rightarrow$ child to be missing when the child is reachable via the sibling pointer of its predecessor.



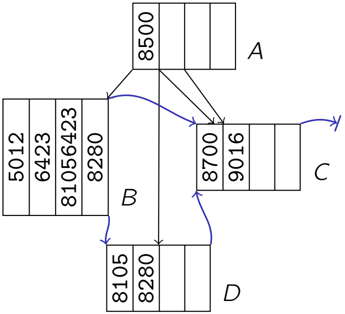
[20]Lehman and Yao. *Efficient Locking for Concurrent Operations on B-Trees*, TODS 6(4), 1981.

B^{link}-Trees

With the relaxation, node splitting and parent updates can be separated.

```
1  latch & read page $B$ ;
2  create new page $D$ and latch it
     ;
3  populate page $D$ ;
4  set next pointer $D \rightarrow C$ ;
5  un-latch $D$ ;
6  set next pointer $B \rightarrow D$ ;
7  adjust content of $B$ ;
8  un-latch $B$ ;

9  latch & read $A$ ;
10 adjust content of $A$ ;
11 un-latch $A$ ;
```



→ Lines 9–11 can be deferred to a later time.

B^{link}-Trees

With the relaxation stated before, lines 1–8 already represent a correct B^{link}-tree.

- Lines 9–11 are, in a sense, only applied for performance reasons.

The parent could be updated also at a **later time**:

- As a “clean-up process” triggered when the update has completed.
- When the next search traverses the tree.
- During database maintenance.

In fact, even the page latches can be avoided when pointer updates and record deletions can be done **atomically**.

PostgreSQL, *e.g.*, uses B^{link}-Trees.

Optimistic Concurrency Control

- So far we've been rather **pessimistic**:
  - we've assumed the worst and prevented that from happening.
- In practice, conflict situations are not that frequent.
- **Optimistic concurrency control:** Hope for the best and only act in case of conflicts.

Optimistic Concurrency Control

Handle transactions in **three phases**:

1. **Read Phase.** Execute transaction, but do **not** write data back to disk immediately. Instead, collect updates in a **private workspace**.
2. **Validation Phase.** When the transaction wants to **commit**, test whether its execution was correct. If it is not, **abort** the transaction.
3. **Write Phase.** Transfer data from private workspace into database.

Validating Transactions

Validation is typically implemented by looking at transactions'

- **Read Sets** $RS(T_i)$**:** (attributes read by transaction T_i) and
- **Write Sets** $WS(T_i)$**:** (attributes written by transaction T_i).

backward-oriented optimistic concurrency control (BOCC):

Compare T against all **committed** transactions T_c.
Check **succeeds** if

$$T_c \text{ committed before } T \text{ started} \quad \textbf{or} \quad RS(T) \cap WS(T_c) = \emptyset \ .$$

forward-oriented optimistic concurrency control (FOCC):

Compare T against all **running** transactions T_r.
Check **succeeds** if

$$WS(T) \cap RS(T_r) = \emptyset \ .$$

Multiversion Concurrency Control

Consider the schedule

$$r_1(x), w_1(x), r_2(x), \overset{\color{red}{t}}{\downarrow} w_2(y), r_1(y), w_1(z) \quad .$$

✎ **Is this schedule serializable?**

- Now suppose when T_1 wants to read y, we'd still have the "old" value of y, valid at time t, around.
- We could then create a history equivalent to

$$r_1(x), w_1(x), r_2(x), r_1(y), w_2(y), w_1(z) \quad ,$$

which is **serializable**.

MVCC in Practice—Read-Only MVCC

A simple form of MVCC is the **Read-Only MVCC**:

- **Read/write transactions** use concurrency control as before (*e.g.*, 2PL)
- **Read-only transactions** do not acquire any locks. For each read operation $r(x)$ of a read-only transaction T_{RO}, read the version of x that existed when T_{RO} started.

That is, read-only transactions see a **snapshot** of the database as of the time when they started.

Problem:

- Must mark each data object with **commit time** of transaction.

“Read-Consistency” MVCC

Oracle implements “read committed” ($\nearrow$ slide 336) using the **“Read-Consistency” protocol**:

- **read-only transactions** are treated as in the Read-Only protocol.
- **writes in read/write transactions** acquire long-duration write locks.
- **reads in read/write transactions** do not acquire read locks; they read the most recent version of any data object.

$\rightarrow$ Reads only return committed values ($\leadsto$ read committed).

$\rightarrow$ Read-only transactions see consistent state (unlike in read committed).

$\rightarrow$ Readers never block writers and vice versa.

Snapshot Isolation

A modification of the same idea yields **snapshot isolation**.

- All **reads** of any transaction T see the version that was current when T started.
- All **writes** must satisfy the **"first committer wins"** property. A transaction T is allowed to commit only if there is no other transaction T' such that
  - (a T' committed between the start and commit time of T **and**
  - (b T' updated a data object that T also updated.

  Otherwise, T aborts.

To test "first committer wins," compare write sets of T and T'.

- Snapshot isolation is implemented, *e.g.*, in Oracle, SQL Server, PostgreSQL

Wrap-Up

ACID and Serializability

To prevent from different types of **anomalies**, DBMSs guarantee **ACID properties**. **Serializability** is a sufficient criterion to guarantee **isolation**.

Two-Phase Locking

Two-phase locking is a practicable technique to guarantee serializability. Most systems implement **strict 2PL**. SQL 92 allows explicit **relaxation** of the ACID isolation constraints in the interest of performance.

Concurrency in B-trees

Specialized protocols exist for concurrency control in B-trees (the root would be a locking bottleneck otherwise).