Deriving Differential Auxiliary Information for Self-maintainable Views

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Summary
We present an innovative method of deriving an auxiliary view from the running databases. During view maintenance, our method minimized to accesses base relations as well as not to re-execute the view definition again. Using this approach, long transactions from the materialized view maintenance intermingled with database transactions can be committed without fear of abort. We use a formal algebraic approach and develop relevant theorems and proofs. For comparison purposes, four corresponding methods are compared: the bottom line base relation method (Base), incremental base relation method (BaseInc), auxiliary view method (AV), and differential file method (DF). We also consider the handicapped cost of storing information as we examine the worst as well as the best case scenarios of our method. Experimental analyses show that the DF method is superior to the other methods in a large data environment of up to 10 Terabyte tuples of relations. The results show that compared with the other methods, the DF method can considerably reduce the number of IO’s. Our most important finding is that the DF method can successfully update the aggregate SPJ views (practically) self-maintainable in the tuple level as well as (theoretically) is independent of the DBMS.

Key words:
Referential Integrity, View Maintenance, Differential Files, Self-maintainability.

1. Introduction

Views are materialized to provide fast access to information that is usually integrated from several distributed data sources. One of the critical weaknesses of materialized views (MV) is that the views are liable to become outdated or desynchronized with the source data as changes are made to the source data upon which the views are defined. In order to guarantee the correctness (or currency) of the MV, all changes to source data have to be applied to the views. Many studies have been extensively undertaken on what is called the view maintenance problem or the materialized view update [13, 14, 20].

In response to source changes, a view can be either recomputed from the source data or maintained incrementally without accessing the source data (called self-maintenance).

We propose an innovative view update algorithm that does not access base relations. We use a differential file and an Auxiliary Integrated File to fully maintain materialized views without accessing base relations. The DF is the changed portion of a base relation, which is sometimes called the delta of a base relation. The AIF is defined as auxiliary information derived from a base relation through the referential integrity constraints between relevant base relations.

The rest of the paper is organized as follows. Section 2 discusses the related works. Section 3 introduces a formal approach to materialized view maintenance. The formal algebra is in section 4. We present cost functions with the parameters and performance analyses in section 5, and section 6 concludes our paper.

2. Related Works

Self-maintenance is a notion that can be defined as maintaining views by materializing supplementary data so that the view can be maintained without (or at least mostly without) accessing base relations. The notion was originally introduced by Blakeley [2]. The main idea is based on a Boolean expression with sufficient and necessary conditions on the view definition for autonomously computable updates that can be called self-maintainable views. Blakeley’s algorithm is a special case of the counting algorithm applied to select-project-join expressions (no negation, aggregation, or recursion). Theodoratos et al. [20] summarize issues extensively related to self-maintainability, and suggest a view selection approach based on a DAG (Please write out the whole word DAG then put abbreviation in parentheses) method. Several notable articles that deal with self-maintenance aim to develop algorithms related to the integration and the maintenance of information extracted from heterogeneous and autonomous sources [11, 18].

Algebraic approaches for maintaining materialized views are discussed in [2, 19, 7, 11, 9, 20]. Excluding
conventional database approaches, [2]. Quian & Widerhold [19] present an algorithm for incremental view maintenance based on finite differencing techniques (later corrected in [8]). The algorithm derives the minimal incremental changes in an arbitrary relational expression for a view modification by replacing the original relational algebraic expression with an efficient and incremental re-computation. They considered two types of operations: insertions and deletions. However, the algorithm uses source relations and thus it lacks the self-maintenance notion. Griffin & Libkin in [7] extend the techniques in [19]. [11] proposes to include functional dependencies. [9] integrates outer joins. These references do not consider the concepts of referential integrity for the maintenance of materialized views. In this paper, some of the common notations (mainly from [19], [7], and [8]) are extended to present some propagation rules for materialized views based on referential integrity constraints.

There has been some research that considers the database system as a rule system [6, 5]. Widom et al. [5] (enhanced from [6]) present a comprehensive survey on the roles played in materialized views. In that paper, the rule is classified as a constraint or a trigger in that the constraint is descriptive while the trigger is procedural. (However, in this paper we use the term ‘constraint’ interchangeably with ‘rule trigger.’) Though not directly related, there are several works [16, 10, 11] corresponding to this method that have the potential to extend referential integrity constraints to the maintenance of database views. When a referential integrity rule invokes cascade among database rules in the DBMS, [16] presents the run time execution problem and the safeness condition respectively. [15] investigates the view maintenance problem with inclusion dependency but no referential integrity rules.

Entity and referential integrity rules are the most fundamental constraints that any relational database should satisfy [15]. The entity integrity rule starts from the selection of a candidate key and referential integrity starts from the selection of a foreign key. Codd’s definition of referential integrity is ‘No component of a foreign key is allowed to have an l-marked value,’ where an l-marked value means a null value of the type ‘value does not exist,’ or ‘value at present exists but is unknown,’ or ‘value is inapplicable’.

Database rules, including referential integrity constraints, are utilized in maintaining materialized views in several articles such as [17, 18]. Quass et al. [18] and Mohania et al. [17] use the referential integrity constraint to determine whether a base relation is participating in the views, and [18] extends the works of [19] and [7] to transform change propagation equations into more efficient ones. They use an auxiliary view (in [17] ‘auxiliary relation’, in [12] ‘auxiliary data’, and ‘complements’ in [14, 15]) in order to maintain a select-project-join (SPJ) view without accessing base relations at the sources. However, the validity and the performance of these methods are strongly dependent upon query types, as long as the view conditions can screen the corresponding base relation. This is discussed in the motivational examples in Section 2.

In applying referential integrity to view maintenance, the work of Quass et al. [18] (called, ‘AV method’ in this paper) is slightly related to our approach, but there are several differences. While Quass et al. assume that the structure of the referential integrity conditions should be a tree; our approach does not assume it. Thus in the AV method, any kind of cycles in a database schema, including self-join, cannot be supported. For example, a transitive closure algorithm [18] cannot support the schema in Section 2. Another difference is that it can be said that the AV method uses the rule in a macroscopic way (i.e., using the rule to find a corresponding ‘relation’), but our method uses it in a microscopic way (i.e., using it to find a corresponding ‘tuple’ in the relation). This is one of the innovative features of our approach that differentiates it from others.

3. Algebraic Representation

It is assumed that the materialized view in this paper is not from independent relations, but from referentially integrated relations. Thus, a relation has an attribute called a foreign key such that it has a referential integrity condition with the key of some relation. (We will call it a RI condition or a RI constraint.) In that case, one relation is called a referenced relation (e.g., DEPT) and the other a referencing relation (e.g., EMP). The referencing relation can be called a fact table and the referenced relation a dimension table. From this point of view, they are called a star schema or a snowflake schema in the materialized environment. Let s represent a referencing relation and r be a referenced relation respectively throughout this Section. The condition of the RI is typically assumed to have one of the representative RI conditions such as restrict, cascade, and nullify [16]. We have a focus on the RI condition that can also be extended to a modified form or a nested form.

We explain the view updates in terms of the changes to base relations and how these changes affect other relations. They are classified as follows: the insertion in the referenced relation case, the deletion in the referenced relation case, the deletion in the referencing relation case, and the insertion in the referencing relation case, including these changes with base relations. The orders are not significant. The last case is the most complicated one to solve. Thus a new schema called an AIF in terms of referential integrity conditions is presented.
The following notations and algebraic expressions on the data base schema $R_1$ are introduced:

\[ R_1 := \emptyset \]

empty set

\[ \sigma_{\text{condition}}(r) \]

selection with condition $P$

\[ \pi_{A}(r) \]

projection over an attribute $A$

\[ r \bowtie r \]

disjoint union

\[ r \bowtie r \]

contained difference

\[ r \bowtie r \]

natural join

**Definition 1** Let a tuple $t$ defined on a relation $r_i$ for $i \in \psi$ be represented as $t(r_i)$. An inserted tuple $\Delta t(r_i)$ and a deleted tuple $\nabla t(r_i)$ belong to the insertion $\Delta r_i$ and the deletion $\nabla r_i$ of the relation $r_i$, respectively. It can be specified by a key $k$ and by a foreign key $(fk)$ as $t(r_i, key)$ and $t(r_i, fk)$ respectively.

**Definition 2** A differential file $DF(r_i)$ of a base relation $r_i = \{key, a_{i1}, a_{i2}, \ldots, a_{im}\}$ can be defined as $\{key, a_{i1}, a_{i2}, \ldots, a_{im}, \text{operation}, sysdate\}$ for $i, m \in \psi$, where operation is an operation type having two possible values i.e. ‘insert’, ‘delete’. A modification is a deletion and an insertion in series having the same sysdate. The sysdate is the timestamp recorded by a committed transaction to the base relation $r_i$. Then the differential file consists of insertions and deletions represented algebraically as $DF(r_i) = \Delta r_i \bowtie \nabla r_i$. □

**Example 4.1** Let’s consider the instance of relations $Pt$, $Su$, and $Sp$ defined as those in Section 2. The corresponding differential files of $Pt$ and of $Sp$ (i.e., $DF(Pt)$ and $DF(Sp)$) are respectively as shown below. There are no changes in the other tables. The first tuple of $DF(Pt)$, say $\{P4, printer, red, 300, delete, 10/08/06\}$, represents that a ‘printer’ with price ‘300’ supplied by the manufacturer ‘red’ was deleted at time ‘10/08/06’. Similarly, in the differential file of $Sp$ $DF(Sp)$, we can tell that the first and the second tuples indicate that the product ‘P1’ supplied by ‘S1’ was modified at time ‘05/08/06’. Then a product ‘P6’ supplied by ‘S4’ was inserted.

<table>
<thead>
<tr>
<th>$pk$</th>
<th>$pn$</th>
<th>$mfr$</th>
<th>$price$</th>
<th>$operation$</th>
<th>$sysdate$</th>
</tr>
</thead>
<tbody>
<tr>
<td>P4</td>
<td>printer</td>
<td>red</td>
<td>300</td>
<td>delete</td>
<td>10/08/06</td>
</tr>
</tbody>
</table>

**Table 4.1**: Differential file $DF(Sp)$

**Table 4.2**: Differential file $DF(Pt)$

**Corollary 3** A new base relation (after image) $r_i^{\text{new}}$ can be expressed as the base relation (old image) $r_i$ and its $DF(r_i)$ as follows: $r_i^{\text{new}} = r_i \bowtie DF(r_i) = r_i \bowtie (\Delta r_i \bowtie \nabla r_i) = (r_i \bowtie \Delta r_i \bowtie \nabla r_i)$ for $i \in \psi$. □

Notice that the schema adjustment between a base relation and the corresponding DF is assumed to be exchangeable in this paper. Since the DF has two more columns (i.e., operation and sysdate) than the corresponding base relations. For simplicity, we do not create an additional operator to adjust between them. In this paper, we follow the notation used in previous research [7, 8, 18, 19].

**Example 4.2** The new relation $Pt^{\text{new}}$ is derived from $Pt$ (in example 2.1) and with $DF(Pt)$ in Table 4.1 as:

$Pt^{\text{new}} = Pt \bowtie DF(Pt) = \{(P1, computer, red, 2000), (P6, DVD, yellow, 2000), (P7, computer, green, 3000)\}$

**Definition 4** A relation $r_i$ satisfies a RI condition, which can be represented as: $r_i^\text{RIcondition}$, where the superscripted [condition] can optionally represent a RI condition detail.

For example, an insertion in $s$ due to an RI condition of a deletion in a relation $r$ can be represented: $\nabla s \bowtie r^\text{RIcondition}$. □

**Example 4.3** Note that there is an RI condition between the relations $Cu$ and $Or$ in section 2. If the RI condition is ‘On delete Cascade’ and a deletion in $Cu$ (e.g., ‘C2’) may affect the relation $Or$, then the corresponding tuple in $Or$, (e.g., $O2, C2, 5, 2500$) should also be deleted. It can be represented as follows:

\[ \nabla Or^{\text{RIcondition}}Cu \] = \{(O2, C2, 5, 2500, delete, 05/08/06)\}

A change in a relation can cause other changes in turn due to the referential integrity constraints. The nested changes of base relation can be represented in general as follows: In that case, the changes can be represented as a nested form in terms of RI conditions.

**Corollary 5** A change in the referenced relation ($r_i$) can generate a change due to the RI condition to the corresponding referencing relation ($r_s$) that results, in turn, in a change recursively to $r_n$ and finally it may effect a change in the relation $s$. The generalized deletion and insertion due to RI condition in $s$ can be represented respectively:

\[ \nabla_s^{\text{RIcondition}}Cu \] and \[ \Delta_s^{\text{RIcondition}}Cu \]

\[ \nabla_s^{\text{RIcondition}}Cu \] and \[ \Delta_s^{\text{RIcondition}}Cu \]

**Example 4.4** Note that there also is an RI condition between the relations $Or$ and $Sp$ in section 2. In addition to the above example, a deletion in $Or$ will also affect relation $Sp$, thus the corresponding tuple, i.e., (P1, S4, O2,
5, 2500) should also be deleted. Therefore in order to
(succcessfully) delete a tuple in Cu, two more deletes in Or
and in Sp are needed: \( \nabla Sp_{RI}^t \times (\nabla Or \cup Cu) \) = \{ (P1, S4, O2, 5, 2500, delete, 05/08/06) \}

Note that by the RI conditions, the cascading RI
representations of a deletion and an insertion are one of the
DF tuples of the corresponding relation:

\[
\Delta s_{RI}^\nabla = \nabla r_{RI} \quad \text{and} \quad \Delta s_{RI} = \nabla r_{RI} \nabla s_{RI}
\]

For example, there are the two more deletes in the above
equations (4.3 and 4.4) that did not originate from their
own transactions, the deleted tuples are included among
the DFs of their own, i.e., DF(Or) and DF(Sp) respectively.

**Theorem 6**

\[
\nabla s_{RI} = \nabla r_{RI} \nabla s_{RI} = \nabla r_{RI} \nabla s_{RI} = \nabla r_{RI} \nabla s_{RI}
\]

Proof: If there is no change in a referencing relation s, i.e.,
\( \Delta s = \nabla s = \emptyset \), then the above equations trivially hold. If
there exists an insert transaction in s, then it will be committed
if the relevant key exists in the referenced relation r, unless the
transaction will be aborted. In the two cases, there is no
change due to RI in the relation r. If there is a delete
in r, then the transaction commits without an
inquiry to r. Therefore no changes happen in r due to the
RI by the changes of s. □

From the above theorem, we can get the following trivial result.

**Corollary 7**

\( \nabla s_{RI} \ni r \rightarrow \nabla s_{RI} \)

**Theorem 8**

A deletion in r triggers a change in s by the RI
condition. Then the following rule holds:

\[
(s \backslash (\nabla s_{RI} \nabla r_{RI}) \ni \nabla s_{RI} \cup \nabla r_{RI}) \ni \nabla r \rightarrow \emptyset
\]

Proof: If there is no deletion in the referenced relation r, then
the above holds. Suppose that there exists a deletion but there
exists no insertion by RI or no deletion by RI in a
referencing relation s. Further suppose that there exist some
tuples in the referencing relation corresponding to the deleted
tuple of the referenced relation. This is a referential integrity
violation. It contradicts the above assumption. Thus, as long
as there remains a foreign key in a referencing relation, there
will be two cases with respect to the RI condition: (1) If the
RI condition is ‘restrict’, then the deletion in the referenced
relation cannot be committed, (which means \( \nabla r = \emptyset \)). (2) If
the RI condition is ‘cascade’ or ‘nullify’, then the changes
caused by the RI will occur in the referencing relation s (say, \( \Delta s_{RI} \uplus \nabla s_{RI} \)). This represents that the above rule holds. □

**Example 4.4**

\( (Or \ni \nabla Or_{RI} \nabla Cu) \ni \nabla Cu = \emptyset \)

**Theorem 9**

\( s \ni \nabla r = \emptyset \)

Proof: For some tuples \( t \) such that \( t(\nabla r) = s_{fk} \), if \( \exists t \) such as \( \exists t, t(\nabla r) = t(\nabla r) \), this violates uniqueness of key. □

**Theorem 10**

\( s \ni \nabla r \rightarrow \nabla s_{RI} \ni \nabla r \)

Proof: \( s \ni \nabla r \rightarrow (s_{new} \ni \nabla s \uplus \nabla s) \ni \nabla r \rightarrow \nabla s_{new} \ni \nabla r \). The proof is based on Corollary 3, and
Theorem 4.2 respectively. □

Therefore, for all the RI conditions, the join of the deleted
tuple(s) (i.e., \( \nabla r \)) in the referenced relation with the tuples in
referencing relation s should be Null or equivalent to the deltas.

**Example 4.5**

Suppose that a product ‘P6’ is not to be
deployed in the previous example, i.e., a delete transaction for the tuple \{P6, DVD, yellow, 20\} is issued at time
01/09/06. The transaction will then be executed in terms of the RI conditions:

1) If the RI condition is ‘On delete restrict’ on the key of Pt,
i.e., \( t(\nabla Pt) = \emptyset \) (restrict), the transaction will be aborted
due to RI in Sp.

2) If the RI condition is ‘On delete cascade’ on the key of Pt,
i.e., \( t(\nabla Pt) = \emptyset \) (cascade), it will delete a tuple in Sp (i.e.,
\{P6, S1, O3, 20, 250\} will be deleted). The last case ‘On delete
nullify’ on the key of Pt, i.e., \( t(\nabla Pt) = \emptyset \) (nullify), it will modify another tuple in Sp as \{P6, S1, O3, 20, 250, delete, 01/09/06\} and \{NULL, S1, O3, 20, 250, insert, 01/09/06\} are appended to DF(\( s_{RI} \)) in series.

**Definition 11**

An RI check condition is a Boolean function that
there is an insertion transaction in the referencing
relation s, and there follows an RI check to see if the
foreign key in s corresponds to the key of the referenced
relation r. The RI check condition is:

\( t_r(\nabla s_{fk}(RI\text{check})) \) is true, if \( \exists t_r, t_r(\nabla s_{fk}) = t_r(\nabla r) \), otherwise false.
Lemma 12 An insert transaction in the referencing relation is committed, the RI check condition corresponding to the insertion is true.

Proof: Suppose the RI check condition is false but an insertion transaction in the referencing relation is committed. The RI check condition is false, \( \triangle s.fk \neq r.key \). It violates \( f.k \) integrity. \( \Box \)

Example 4.6 There are two RI constraints in the relation \( Cu \). Thus the RI check is fired to check the relevance of the insertion in \( Cu \). When a tuple \( \{C3, BC, R&D, C2\} \) is inserted in \( Cu \), and then the relevant two RI constraints are fired. One is to acknowledge \( Cu.\{BC\} = Ar.\{BC\} \), and \( Cu.\{C2\} = Cu.\{C2\} \). If the results of both RI checks are true, the insertion transaction will be committed. If any of the two RI checks are not true, the transaction will be aborted.

Example 4.8 If there is a committed insertion \( \{P6, S4, O3, 10, 350\} \) in \( Sp \), then three RI check tuples are appended (one each) to the RI differential file of \( Pt \), that of \( Su \), and that of \( Or \). A tuple \( \{P6, DVD, yellow, 200\} \) is appended to \( AIF \) called \( AIF(Pt) \), a tuple \( \{S4, A, CA\} \) to \( AIF(Su) \) and a tuple \( \{O3, C1, 30, 300\} \) to \( AIF(Or) \).

The descriptions of the equations are the propagation rules slightly modified from the rules based on [19, 7, 8].

5. View Maintenance

The view self-maintainability (SM) is one of the important considerations in this paper, defined as maintaining a view in response to changes of database relations using only the view and the differential files to the base relations, without accessing the base relations. The SP view satisfies SM [9] (we will show this algebraically in the next section), but joint views might need some auxiliary information from other relations [18]. In order to achieve SM of views, auxiliary relations might be prepared before executing view expressions. In other words, with this, auxiliary views can satisfy SM in view execution time. The auxiliary view [18], auxiliary relation [17], auxiliary data [12], and complements [15] all need access to base relations in making their auxiliary information.

A view is a mapping between a query and data. If a join query is issued, then the join condition is mapped to RI constraints and is substituted with the data. Notice that the data is assumed to synchronize via corresponding DFs of database relations. Duplicated tuples in DFs should be eliminated by some algorithm that all the tuples are removed except the first and the last. The AIF is derived from the corresponding base relation, so the tuples of AIF may be duplicated with tuples within the AIF as well as those of DF of the corresponding relation. Thus the duplicated tuples should be eliminated. Notice that if the delta is refreshed immediately, there will be no duplicated tuples (within the AIF). In general, the duplicated tuples should be minimized in any view update policy (e.g., immediate update, periodic update, and deferred update, etc).

The schema of \( AIF(x_i) \) of a base relation \( z_i = \{k_\psi, a_{i1}, a_{i2}, \ldots, a_{im}, \text{sysdate}\} \) for \( i, \psi \in \psi \), where \( \text{sysdate} \) is the timestamp recorded by the RI check condition if answered true. It is assumed that the timestamp recorded by the RI check condition is the same as the relevant committed transaction to the base relation \( r_i \). So the AIF is appended by the order of \( \text{sysdate} \).

Given an RI (e.g., \( z_s.fk \subseteq x_i.\text{key} \)) and view refresh time, the algorithm for duplicate elimination on the \( AIF(x_i) \) is:

1. at first within the \( AIF(x_i) \) and then \( DF(x_i) \).

Algorithm 1: Duplicate Elimination of AIF

Input: 
- \( DF(r_i) \) for \( i \in \psi \)
- Output: consistent \( DF(r) \)

Procedural Steps:

1. \( t_1 \leftarrow \text{last refresh time}; \; t_2 \leftarrow \text{current refresh time}; \; s \leftarrow m \)
2. (Duplicate elimination within AIF) Get next tuple of \( AIF(r_i) \) with \( t_1 < AIF(r_i).\text{sysdate} \leq t_2 \).
   - If no tuple found, go to Step 6.
3. Delete \( AIF(r_i).\text{key} \) from the hash index if there exists.
   - Go to Step 2.
   - Else continue.
4. (Duplicate elimination with DF) Get next tuple of \( DF(r_i) \) with \( t_1 < DF(r_i).\text{sysdate} \leq t_2 \).
   - If no tuple found, go to Step 6.
5. Delete \( AIF(r_i).\text{key} \) from the hash index if there exists.
   - Go to Step 4.
   - Else continue.
6. If \( s = 1 \), then stop.
   - Else do: \( s \leftarrow s - 1 \); Go to Step 2.

In this paper, it is assumed that the AIF has index on the candidate key. It is not a hard assumption, for the duplicated tuples in AIF can easily be deleted, even if there is no index. The index may be a hash index or a filter bit vector where it can contain only a key value. In terms of storage requirements, the hash index approach requires approximately \( 20^8 \times 0.7 = 228.57 \) bits per key value, where the size of key is assumed 20 bytes and 70% storage utilization of the hashing scheme, while the bit vector requires one bit per key.
5. Performance Analyses

The following values are assigned to the parameters for the analysis. The block size is generally assumed to be $B = 4000$ bytes, and the I/O cost $C_{I/O} = 25$ ms/block. The cardinalities of the base tables are assumed to be examined from 1,000 and 10 Terabyte tuples respectively, and the size of the differential file is varied in the experiment. The communication speed varies from a very low case to a high-speed case, that is $C_{comm} = 100Kbps$ ~ $10Mbps$. Tuples are filtered from a no screening case ($\alpha_s = 1.0$) and a highly screened case ($\alpha_s = 0.001$). According to the above parameters and cost functions presented in Section 6, the following four methods are analyzed: (1) the base table method ($\text{Base}$), (2) the base table incremental method ($\text{BaseInc}$), (3) the auxiliary view method ($\text{AV}$), and (4) the differential file method ($\text{DF}$).

Fig. 6.1 and Fig. 6.2 show that the total costs of the three methods are strongly dependent on both the selectivity with the screen factor and the communication speed. Fig. 6.1 represents deleted the costs of all methods are decreased along with communication speed, and the cost of $\text{DF}$ is consistently less than 10% of base method in any communication speed. This shows that the size of the data per se is the most critical factor. If the tuples are filtered highly (up to about 0.01), as in Fig 6.2, the Base method and the $\text{BaseInc}$ method are less advantageous than the $\text{DF}$ method or the $\text{AV}$ method. However, if tuples are less screened (so, the screen factor is up to about 1.0), the $\text{DF}$ method is less advantageous than the $\text{AV}$ method and, in some cases, even less advantageous than the Base method.

6. Conclusion and Future Research

In this paper, we have presented an effective method that can derive auxiliary information on which we utilized the potentials of the referential integrity constraint in DBS. Our method, called $\text{DF method}$, uses a $\text{DF}$ (Differential File). The $\text{DF}$ is defined as the changed portion of a base relation, which is extracted from the (referenced) base relation by checking the referential integrity constraint. In developing algebraic expressions, a series of robust definitions and theorems having to do with several rules for auxiliary information are generated. When updating $\text{SPJ}$ aggregate views, the relevant base tables are not re-accessed in order to generate auxiliary relations.

The performance analysis has proven to be useful in answering the following questions: (1) Do the $\text{DF}$ and the $\text{AIF}$ have a major role in the view maintenance? (2) What will be the mode of cost trajectories along with the data increasing (i.e., diverging or converging)? The performance analysis indicates that the total cost of the $\text{DF method}$ is closely dependent on the size of the $\text{DF}$ and the $\text{AIF}$, the selectivity (screen) factor, and the transmission rate factor. As these three factors increase with low transmission rate, so the total cost does. When the data size became increasingly large, the cost ratio of the $\text{Base}$
and BaseInc method diverged too much, but that of the DF method increased slightly and that of the AV method was intermediate. However, cost benefits are negligible in a worst-case scenario with such factors as a huge DF and AIF (up to the size of the base tables), and no selection cases. The DF method is still superior to the Base and BaseInc method as well as to the AV method in that it can make the SPJ aggregate views independent of the current database relations. By using this approach, long transactions from the materialized views intermingled with conventional database transactions can be committed without fear of transaction abort. In the worst-case environment the experimental results represent that managing the referencing relation and the view is important.

Our future research will be extended to various issues as follows. A data warehouse setting with various views will be developed in terms of DFs. View selection with multiple views and concurrency control issues using these can be worth tackling to relieve various distributed replica settings. In view of this, a view maintenance engine can be implemented. Our method has the potential to deal with the approximation problem of data streaming [1], accuracy of continuous queries and group queries [12], web applications [5], and mobile settings, etc., and therefore it can be easily extended to a myriad of environments.

References

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