Multiple-View Self-Maintenance in Data Warehousing Environments

Nam Huyn [*]
Stanford University
huyn@cs.stanford.edu

Abstract

A data warehouse materializes views derived from data that may not reside at the warehouse. Maintaining these views efficiently in response to base updates is difficult, since it may involve querying external sources where the base data reside. This paper considers the problem of *view self-maintenance*, where the views are maintained without using all the base data. Without full use of the base data, however, maintaining a view unambiguously is not always possible. Thus, the two critical questions that must be addressed are to determine, in a given situation, whether a view is maintainable, and how to maintain it.

We provide algorithms that answer these questions for a general class of views, and for an important subclass, generate SQL queries that test whether a view is self-maintainable and update the view if it is. We improve significantly on previous work by solving the view self-maintenance problem in the presence of multiple views, with optional access to a subset of the base data, and under arbitrary mixes of insertions and deletions. We provide better insight into the problem by showing that view self-maintainability can be reduced to the problem of deciding query containment.

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1 Introduction

Data warehouses have gained importance in recent years ([RED, IK93, Z*95]). A data warehouse is a collection of materialized views derived from data that may not reside at the warehouse. As a benefit, user queries can often be evaluated much more cheaply using these stored views than using the base relations. The problem, however, is that the views must be updated to reflect changes made to the base relations. While maintaining these views incrementally is often significantly more efficient than recomputing them from scratch (as done in most current data warehouses), it can still be expensive. For instance, in response to an update to a base relation, incremental maintenance of views defined as a join may involve looking up the non-updated base relations, which may reside in external sources.

Thus, in data warehousing environments where maintenance is performed locally at the warehouse, an important incremental view-maintenance issue is how to minimize external base data access. The idea of avoiding base access to speed up view maintenance is illustrated in Figure 1. We take the following approach to reduce maintenance costs. In response to a base update, we try to maintain the views using information that is strictly local to the warehouse. This information includes the view definitions and the contents of all the views. Only when we fail to do so do we resort to accessing the base relations.

As a result of not using all the base relations, there may be situations where there is not enough information to maintain a view unambiguously, even if we are given the specific contents of the views, a subset of the base relations, and the base update. Such situations never arise in traditional work on materialized view maintenance ([GM95, Kuc91, GMS93, SJ96]) where all the base data is usually assumed to be available. Thus, an important question (originally considered in [TB88, Hu96]), which was never raised in traditional view-maintenance work, is to determine whether

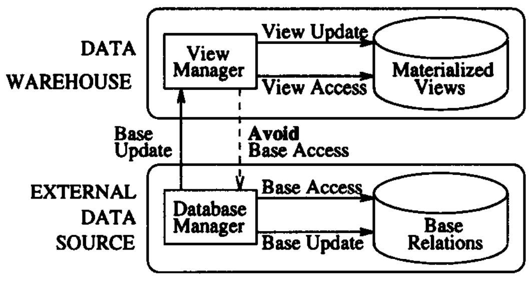

Figure 1: Saving Base Access in View Maintenance.

a view is *maintainable*, that is, guaranteed to have a unique new state, given an update to the base relations, an instance of the views, and an instance of a subset of the base relations. As a shorthand, such a view is said to be *self-maintainable* [1] in the given situation. A second question, the main question in traditional view-maintenance work, is how to bring the view up to date using only the given information. Together, these two questions define the *view self-maintenance* problem.

Previous work on view self-maintenance specific to a given situation ([TB88, GB95, Hu96]), however, only considered the special case where *no other materialized views* and *no base relations* are used to maintain a given view. We call this case *single-view self-maintenance*. Applying these methods to maintain a warehouse that contains several views, i.e., by maintaining the views separately from each other, often fails to maintain the warehouse when actually the views are self-maintainable collectively. The following example illustrates the need to use all the views to maintain a warehouse.

Example 1.1 Consider a data warehouse with two materialized views $V_1(X,Y,Z)$ and $V_2(Y,Z)$, each defined in terms of the base relations $R(X,Y)$, $S(Y,Z)$, and $T(Z)$ as follows:

```
CREATE VIEW V1(X,Y,Z) AS
SELECT R.X, R.Y, S.Z FROM R,S,T
WHERE R.Y = S.Y AND S.Z = T.Z

CREATE VIEW V2(Y,Z) AS SELECT * FROM S
```

That is, V_1 is the natural join of R, S, and T, while V_2 is a copy of S. Suppose we would like to maintain the warehouse in response to the insertion of tuple (a, b) into R, without using either R, S, or T.

First, consider the view instance where $V_1 = \{(a_1, b_1, c_1)\}$ and $V_2 = \{(b_1, c_1), (b, c_2)\}$. While we can infer the contents of S, since V_2 is just a copy of it, we cannot determine the contents of R and T exactly. In fact, it could be that $R = \{(a_1, b_1)\}$ and $T = \{(c_1)\}$, in which case view V_1 is not affected by the insertion of (a, b) into R. But it could also be that $R = \{(a_1, b_1)\}$ and $T = \{(c_1), (c_2)\}$, in which case (a, b, c_2) must be added to view V_1 to keep it consistent with the base relations. Thus, we cannot unambiguously maintain view V_1. V_1 is not self-maintainable under the insertion in this view instance.

Consider another instance where $V_1 = \{(a_1, b_1, c_1), (a_1, b_1, c_2)\}$ and $V_2 = \{(b_1, c_1), (b, c_2), (b_1, c_2)\}$. This time, however, we can infer enough about T to be able to precisely determine the effect of the insertion on V_1. In fact, to evaluate the effect of the insertion on view V_1, we look for tuples from S and T that join with the new tuple (a, b) from R. On the one hand, only one tuple from S qualifies: (b, c_2). On the other hand, to explain the presence of (a_1, b_1, c_2) in V_1, it must be case that T contains (c_2). Thus, we know exactly how to maintain V_1 without even looking at the base relations: add (a, b, c_2) to V_1. While V_1 is clearly self-maintainable under the insertion of (a, b) into R in this view instance, ([TB88, Hu96]) would fail to detect this situation because they attempt to maintain V_1 in isolation from V_2. If V_2 were not available, they would have concluded correctly that V_1 cannot be unambiguously maintained, since the following two base instances, while both consistent with V_1, derive different states of V_1 after the insertion of $r(a, b)$: $R = \{(a_1, b_1\}$, $S = \{(b_1, c_1), (b_1, c_2), (b, c_2)\}$, and $T = \{(c_1), (c_2)\}$ on the one hand, and $R = \{(a_1, b_1\}$, $S = \{(b_1, c_1), (b_1, c_2)\}$, and $T = \{(c_1), (c_2)\}$ on the other hand. But in light of V_2, the latter base instance is clearly not possible. ∎

Thus, to maximize the chance of maintaining the views successfully, we must take full advantage of all the information available, namely, not only the contents of the view to maintain, but also the contents of all the other views. Further, if a base update consists of a set of individual updates to several base relations, it is very important to consider the set of updates as a whole instead of considering each individual update separately. In fact, there are situations where the former approach succeeds to maintain a view but the latter fails. For instance, consider a view that computes whether or not two base relations (with identical schemas) have a tuple in common. Consider a situation where the view is empty and where the same tuple is to be inserted into *both* relations. Clearly, the new state of the view can only be "true". But if we consider the two insertions independently, we cannot unambiguously update the view.

Our work focuses on the *multiple-view self-*

[1]The terms "self-maintainable" and "self-maintenance" have been used in the literature with quite different meanings, depending on the amount of information available. We will more precisely define our notion of self-maintainability later.

maintenance problem where again, the two critical questions are maintainability and maintenance of a view as a function of a given base update, a given instance of all the views, and a given instance of a particular subset of the base relations. Our work is mainly motivated by the desire to speed up view maintenance in WHIPS ([H*95]), a data warehousing system prototype developed at Stanford University that performs on-line update to views. Also, view self-maintenance is an effective approach to avoid asynchronous update anomalies ([Z*95]) inherent in any data warehousing environment, by limiting access to base relations that do no change, for instance. Finally, the self-maintenance approach is also important in any environment where source access is expensive for a variety of reasons: the required base data may be archived; the data source may be temporarily disconnected or even permanently destroyed.

The contributions of this paper are as follows:

- We solve the view self-maintenance problem in the presence of a wide variety of information that is available in practice: multiple views, partial access to the base relations, partial base copies, arbitrary mixes of base insertions and deletions.
- We provide new insight into the view self-maintenance problem by showing that self-maintainability can be reduced to the problem of deciding query containment.
- We show that for conjunctive-query views with no projection, not only efficient (polynomial-time) solutions to the view self-maintenance problem exist but also they can be generated at view-definition time in the form of SQL queries.

Related Work on Self-Maintenance

Self-maintenance generally refers to the problem of maintaining views without full use of the base relations. [GM95] gave an excellent classification of different notions of self-maintenance, based on what information is available for view maintenance. A major distinction is what we call ***compile-time*** vs. ***runtime***. Even though both approaches share the goal of maintaining a view using only the information given (namely the instance of the views, the update instance, and perhaps the instance of some subset of the base relations), they differ in the way they guarantee a view can be maintained. In the compile-time approach, this guarantee is made independently of the instance of the views, the instance of the base relations, and the instance of some update type. "Compile-time" alludes to the fact that these instances are not known at compile-time, but only the view definitions and the update type. In the runtime approach, maintainability of a view is guaranteed on an instance basis: for a particular instance of the views, a particular instance of a subset of the base relations, and a particular update instance. Note that this approach is the more aggressive one, since it may succeed in maintaining a view where the compile-time approach fails. In this paper, we take the runtime approach to self-maintenance, even though "runtime" is not explicitly mentioned.

Within the realm of runtime self-maintenance, we are not aware of any work that addresses the question of self-maintainability with respect to an instance of more than one view or base relation, or under an arbitrary mix of insertions and deletions to the base relations. [TB88] and more recently [GB95] gave self-maintainability conditions (they called conditions for *Autonomously Computable Updates*) for views that are SPJ queries with no self-joins and for updates that are either insertions or deletions to a single base relation. Their method cannot be extended easily to take advantage of multiple views or the base relations, or to handle updates that mix insertions with deletions. [Hu96] addressed the single-view strict self-maintenance problem and solved it more efficiently than [GB95], but only for views that are SPJ queries with no self-joins and for updates that are single insertions. Again, their method cannot be generalized easily. In the realm of compile-time self-maintenance, [GJM96] addressed the single-view self-maintenance for views defined as SPJ queries and under either insertions, deletions, or updates. More recently, [Q*96] solved a different but related problem, namely that of *making* a view self-maintainable by introducing a minimal set of auxiliary views to materialize. However, how to make more than one view self-maintainable was not addressed.

Outline of the Paper

In Section 2, we define the multiple-view self-maintenance problem. Sections 3 through 5 deal with strict self-maintenance. Sections 3 and 4 consider the special subclass of SPJ views with no projection. Section 3 shows an algorithm that generates the view maintenance queries for a self-maintainable view. Section 4 shows an algorithm that generates queries that test self-maintainability, based on reducing self-maintainability to a query-containment problem. Section 5 shows how to extend the previous results to other classes of views. In Section 6, we show how to extend the results for strict self-maintenance to generalized self-maintenance, which can be used to define a strategy for efficient warehouse maintenance. The paper concludes in Section 7. Throughout this paper, results are stated without their proof, which can be found in [Hu97].

2 Defining the Multiple-View Self-Maintenance Problem

Throughout this paper, the warehouse consists of materialized views $V_1, V_2, \ldots, V_m$ derived from base relations $R_1, R_2, \ldots, R_n$. This collection of base relations is referred to as database D. Each V_i is defined by a query Q_i over database D, written as $V_i = Q_i(D)$. A database D is said to be ***consistent*** with view V_i if $Q_i(D) = V_i$. We assume the existence of a database consistent with all the given views but whose content is not known a priori. We use U to denote a ground update to the base relations and $U(D)$ to denote the updated database. We model U as δR_1^-, δR_1^+, δR_2^-, δR_2^+, $\ldots$, δR_n^-, δR_n^+, where δR_j^- (resp. δR_j^+) is the set of tuples to be deleted from (resp. inserted to) relation R_j. Updates are assumed to be self-consistent, i.e., δR_j^- and δR_j^+ have no tuples in common for any j.

Strict and Generalized Self-Maintenance

In strict self-maintenance, no base relations are used for maintaining a view. Thus, given an instance of $V_1, \ldots, V_m$ and an update instance U, view V_k is said to be ***self-maintainable*** under U if its new state (i.e., the state of V_k that is consistent with the updated database) does not depend on the database, as long as the database is consistent with ***all the views*** prior to the update. That is, for every pair D and D':

$$[\bigwedge_{i=1}^{m} Q_i(D)=Q_i(D')=V_i] \Rightarrow Q_k(U(D))=Q_k(U(D')) \quad (1)$$

Self-maintainability is a function of U and $V_1, \ldots, V_m$ (it is also a function of the view definitions Q_i, but that is understood). Note the requirement that D and D' be consistent with all the views, and not just the view to maintain as in single-view self-maintainability. Only when V_k is self-maintainable does it make sense to maintain it, and a maintenance expression is a function of U and $V_1, \ldots, V_m$ (not just V_k as in single-view self-maintenance).

We generalize strict self-maintenance by also allowing access to some of the base relations. Thus, in generalized self-maintenance, given an instance of $V_1, \ldots, V_m$, an update instance U, and the instance of a subset S of the base relations, view V_k is said to be ***self-maintainable*** (under U and with respect to S) if its new state does not depend on the database, as long as the database is consistent with all the views and ***the given base relation instances in S*** prior to the update. In generalized self-maintenance, both self-maintainability and maintenance expression are a function of U, $V_1, \ldots, V_m$, and the base relations in S.

Notation

In this work, we assume a relational database framework in which views are defined by relational queries over base relations. Set semantics is also assumed. Thus, the answer to a query is a set of tuples. We will use the notation of Datalog (ref. [Ull89]) for all the queries involved in our algorithms. This choice is by convenience, even though any other relational languages could be used. Thus, the view definition for view V_1 from Example 1.1 is written as a Datalog query Q_1 with the single rule:

$$v_1(X,Y,Z) \text{ :- } r(X,Y) \ \& \ s(Y,Z) \ \& \ t(Z)$$

where $v_1(X,Y,Z)$ is called the rule's *head*, and $r(X,Y)$, $s(Y,Z)$, and $t(Z)$ the rule's ***subgoals***. By convention, relation names are written in upper case (e.g., V_1, S, and δR^-) and their predicate in lower case (e.g., v_1, s, and δr^-). The ***extension*** of a predicate is the instance of the relation for the predicate. In general, a predicate is called an ***IDB predicate*** if it appears in the head of some rule, an ***EDB predicate*** otherwise. A particular IDB predicate that is used to return the answers to the query is called the ***query predicate***. Thus, in query Q_1, predicates r, s, and t are the EDB predicates, and v_1 the query predicate. Q_1 is an example of a Datalog query with only one rule whose body contains only EDB subgoals. Such a query is called a ***conjunctive query*** (abbreviated CQ) or an SPJ query with only equality comparisons.

For the most part (except Section 5), the queries Q_i that define the views in the warehouse (in terms of the base relations R_j) are assumed to be conjunctive. In rule notation, we write Q_i as H_i :- G_{i1} & $\ldots$ & G_{in_i}, where the head H_i uses predicate v_i for view V_i and each subgoal G_{ij} uses predicate r_l for some relation among $R_1, R_2, \ldots, R_n$. Constant symbols may appear anywhere in a rule. We also assume that the variables in Q_i's body all appear in Q_i's head. Such a query is said to ***have no projection***. For a description of other classes of Datalog queries, see [Ull89].

Query Containment

The main technique used in solving (1) is based on reducing it to an implication problem known in the literature as the query containment (abbrev. QC) problem (see [Ull89]). Given two Datalog queries P and Q that use EDB relations $E_1, \ldots, E_n$ as input, we say that $P \subseteq Q$ if the answer to P is a subset of the answer to Q, for every instance of $E_1, \ldots, E_n$. ***Instance-specific*** QC is a variation of the QC problem where the instance of some of the input EDB relations is fixed. Given two queries P and Q using EDB relations $E_1, \ldots, E_n, F_1, \ldots, F_m$ as input, and given an instance of $F_1, \ldots, F_m$, we say that $P \subseteq_{F_1,\ldots,F_m} Q$ if the answer to P is a subset of the answer to Q for all instances of $E_1, \ldots, E_n$. The EDB predicates f_i, whose extension is fixed, are called *constant predicates*. The EDB predicates e_i are called *variable predicates*. When the extension of the constant predicates is known, we can always reformulate an instance-specific QC problem to a QC problem by eliminating any constant predicate f as follows: replace any subgoal $\neg f(\bar{X})$ with $\bigwedge_{\bar{x}}(\bar{X} \neq \bar{x})$ and any subgoal $f(\bar{X})$ with $\bigvee_{\bar{x}}(\bar{X} = \bar{x})$, where $\bar{x}$ ranges over the tuples in f's extension. However, when the extension of the constant predicates is not known, we would like to find a condition on these predicates that expresses $P \subseteq_{F_1,\ldots,F_m} Q$. Whether or not such condition always exists is still an open question. In the rest of this paper, we will simply use $P \subseteq Q$ to denote $P \subseteq_{F_1,\ldots,F_m} Q$, as it will be clear from the context which input predicate is constant.

3 Generating Queries to Maintain the Views

In this section, we address the question of how to bring a view up to date *if the view is known to be self-maintainable.* Note that if a view is not self-maintainable, there is no unambiguous way to maintain the view correctly without using additional information, such as querying some of the base relations (see Section 6). However, if the view is self-maintainable, we do not need to know what the actual database really is to maintain the view, since we can use any database that is consistent with all the views to propagate the update to the view. But how can we find such a database? The answer lies in the canonical database. Note that the canonical database is defined relative to an instance of the views.

Definition 3.1 *Canonical Database:* Let $V_1, \ldots, V_m$ be views, and for $i = 1, \ldots, m$, let Q_i be a CQ with no projection that defines V_i over relations $R_1, \ldots, R_n$. The canonical database, denoted $\hat{D}$, consists of all the tuples obtained as follows: for each view V_i, every tuple in V_i that matches Q_i's head provides a substitution that grounds all the atoms in Q_i's body; include all these ground atoms in $\hat{D}$. ∎

Example 3.1 Consider views V_1 and V_2 defined by:

$$v_1(X,Y,Z) \text{ :- } r(X,Y) \ \& \ s(Y,Z) \ \& \ t(Z)$$
$$v_2(Y,Z) \text{ :- } s(Y,Z)$$

Suppose $V_1 = \{(a_1,b_1,c_1),(a_1,b_1,c_2)\}$ and $V_2 = \{(b_1,c_1),(b,c_2),(b_1,c_2)\}$. The canonical database $\hat{D}$ in this view instance consists of $R = \{(a_1,b_1)\}$, $S = \{(b_1,c_1),(b_1,c_2),(b,c_2)\}$, and $T = \{(c_1),(c_2)\}$. ∎

Intuitively, we are trying to reconstruct the base relations minimally from all the given views. When each Q_i has no projection, there is a unique minimal reconstruction, which is the canonical database $\hat{D}$. The following lemma states the key property of $\hat{D}$ that allows us to use it to maintain the views.

Lemma 3.1 *$\hat{D}$ is consistent with all the views.* ∎

The following example illustrates how to maintain the views using the canonical database.

Example 3.2 Continuing from Example 3.1, now consider inserting (a,b) to relation R. If V_1 is self-maintainable under the insertion (and with respect to the given view instance), we know we can obtain the same result for the new state of V_1 no matter which database we use to propagate the insertion and that is consistent with the views. We can use $\hat{D}$ in particular. So to compute the tuples gained by V_1, we simply join $r(a,b)$ with $S = \{(b_1,c_1),(b_1,c_2),(b,c_2)\}$ and $T = \{(c_1),(c_2)\}$ to obtain (a,b,c_2). ∎

The following theorem formalizes the use of $\hat{D}$ to reduce the problem of maintaining a view without using any base relation to a view maintenance problem with unrestricted use of the base relations.

Theorem 3.1 *Let $V_1, \ldots, V_m$ be views, and for $i = 1, \ldots, m$, let Q_i be a CQ with no projection that defines V_i over some database D. Let U be an update to D. If V_k is self-maintainable under U, then the new state for V_k is $Q_k(U(\hat{D}))$, where $\hat{D}$ is the canonical database.* ∎

Theorem 3.1 provides us with the following algorithm that computes the incremental view maintenance expressions.

Algorithm 3.1 *Generate Maintenance Query*

```
Input: $Q_1, \ldots, Q_m$, where each $Q_i$ is a CQ with no
    projection that defines $v_i$ using $r_1, \ldots, r_n$ as input.
Output: Queries for incrementally maintaining $V_k$,
    using $v_1, \ldots, v_m$, $\delta r_1^-, \delta r_1^+, \ldots, \delta r_n^-, \delta r_n^+$ as input.
Method:
  1. For $i = 1, \ldots, m$ and $j = 1, \ldots, n_i$, generate rule
     $(A_{ij})$ : $\hat{G}_{ij}$ :- $H_i$, where $H_i$ is the head of $Q_i$
     and $\hat{G}_{ij}$ is the subgoal $G_{ij}$ in $Q_i$'s body whose
     predicate $r_l$ is replaced by predicate $\hat{r}_l$. These
     rules define the predicates $\hat{r}_1, \ldots, \hat{r}_n$ for the canonical database $\hat{D}$.
  2. Generate queries that incrementally maintain
     $V_k$, using predicates $v_k$, $r_1, \ldots, r_n$, $\delta r_1^-, \delta r_1^+$,
     $\ldots, \delta r_n^-, \delta r_n^+$ as input. Call this set of rules $M$.
  3. Let $\hat{M}$ be obtained from $M$ where every occurrence of $r_j$ is replaced by $\hat{r}_j$, for $j = 1, \ldots, n$.
  4. Return $\hat{M} \cup \{(A_{ij}), i = 1, \ldots, m, j = 1, \ldots n_i\}$.
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∎

Step 1 in Algorithm 3.1 essentially computes the canonical database $\hat{D}$. Step 2 generates queries that incrementally maintain view V_k, i.e., that update V_k to the new state $Q_k(U(D))$ using V_k and all the base relations R_i's (the instance of these base relations is actually taken from the canonical database, which is the purpose of Step 3). Many algorithms exist in the view-maintenance literature ([Kuc91, SJ96]) that can generate queries for incrementally maintaining a view using both the view and all the base relations, for example based on algebraic techniques for differentiating query expressions. Using for instance [SJ96] in Step 2, Algorithm 3.1 generates the queries that compute the required insertions to and deletions from a view, in time linear in the size of the view definitions. The size of these queries is also linear. In practice, if these queries are optimized, we may not need to actually construct the entire canonical database as Step 1 would suggest.

Example 3.3 Consider the view definitions for V_1 and V_2 from Example 3.1 and consider the insertion of $r(a,b)$. Let δv_1^+ be the predicate for the set of net insertions to V_1. Algorithm 3.1 generates the following query for δv_1^+:

$$\begin{aligned} \hat{s}(Y,Z) &\text{ :- } v_1(X,Y,Z) \\ \hat{s}(Y,Z) &\text{ :- } v_2(Y,Z) \\ \hat{t}(Z) &\text{ :- } v_1(X,Y,Z) \\ \delta v_1^+(a,b,Z) &\text{ :- } \hat{s}(b,Z) \ \&\ \hat{t}(Z) \ \&\ \neg v_1(a,b,Z) \end{aligned}$$

which can be simplified further to

$$\delta v_1^+(a,b,Z) \text{ :- } v_1(-,b,Z) \,\&\, v_1(-,-,Z) \,\&\, \neg v_1(a,b,Z)$$
$$\delta v_1^+(a,b,Z) \text{ :- } v_2(b,Z) \ \&\ v_1(-,-,Z) \ \&\ \neg v_1(a,b,Z)$$
▮

If a view is not self-maintainable, applying the maintenance queries generated by Algorithm 3.1 may update the view incorrectly. Thus, before applying them to maintain a view, it is important to make sure the view is self-maintainable. The next section provides a decision method.

4 Generating Queries to Test View Self-Maintainability

To determine whether or not a view is self-maintainable under a given update, we compare the effect of the update on the view under different underlying databases: (1) any database consistent with all the views; (2) the canonical database. The following example illustrates this reduction.

Example 4.1 Consider views V_1 and V_2 as defined in Example 3.1 and consider the insertion of $r(a,b)$. First consider the view instance where $V_1 = \{(a_1,b_1,c_1),(a_1,b_1,c_2)\}$ and $V_2 = \{(b_1,c_1),(b,c_2),(b_1,c_2)\}$. V_1 is self-maintainable in this view instance because inserting $r(a,b)$ into any consistent database exactly causes (a,b,c_2) to be added to V_1, which is precisely the same effect as the insertion into $\hat{D}$ has on V_1 (as determined in Example 3.2). Now consider another view instance where $V_1 = \{(a_1,b_1,c_1)\}$ and $V_2 = \{(b_1,c_1),(b,c_2)\}$. $\hat{D}$ in this case consists of $R = \{(a_1,b_1)\}$, $S = \{(b_1,c_1),(b,c_2)\}$, and $T = \{(c_1)\}$. V_1 is not self-maintainable in this view instance, since the insertion into $\hat{D}$ has no effect on V_1 but there is a consistent database, namely $R = \{(a_1,b_1)\}$, $S = \{(b_1,c_1),(b,c_2)\}$, and $T = \{(c_1),(c_2)\}$, where the insertion of $r(a,b)$ causes V_1 to gain (a,b,c_2). ▮

Thus, self-maintainability of a view under a given update can be characterized completely as the following implication problem: for every database D, if D is consistent with all the views before the update, then D derives the same view as $\hat{D}$ after the update. In an equivalent query-containment formulation, we need to decide *DIFF* $\subseteq$ *INCON*, where *DIFF* is the boolean query that D and $\hat{D}$ derive differently after the update, and *INCON* is the boolean query that D is inconsistent with some view before the update. Due to space limitation, the details of *DIFF* and *INCON* are not shown here but can be found in [Hu97]. Each of these queries is a union of conjunctive queries with negation and $\neq$ comparisons. While containment of such queries can be decided using known algorithms such as [LS93], whether or not it can be decided in time polynomial in the size of the view instance is still an open question, since negation applies to the queries' variable input predicates.

In the following, we give a more refined reduction that avoids this undesirable type of negation, thus allowing more efficient containment checking algorithms to be used and, most importantly, self-maintainability to be decided in polynomial time. The key observation here is that instead of considering all possible underlying databases, we only need to check those that are a superset of $\hat{D}$, simply because a database that does not contain $\hat{D}$ cannot be consistent with the views, as formalized in the following lemma.

Lemma 4.1 *Let $V_1,\ldots,V_m$ be views, and for each $i = 1,\ldots,m$, let Q_i be a CQ with no projection that defines V_i over some database D. If D is consistent with all the views, then D contains the canonical database $\hat{D}$.* ▮

The following theorem formalizes the improved reduction, where $D \cup \hat{D}$ represents an arbitrary superset of $\hat{D}$. The use of "set union" makes sense since a database is a set of tuples. Note that to represent a superset of $\hat{D}$, we did not use an arbitrary database

D subject to the constraint $D \supseteq \hat{D}$, precisely to avoid the undesirable type of negation mentioned above.

Theorem 4.1 *Let $V_1, \ldots, V_m$ be views, and for $i = 1, \ldots, m$, let Q_i be a CQ with no projection that defines V_i over some database D. Let U be an update to D. Then V_k is self-maintainable under U if and only if $Q_k(U(D \cup \hat{D})) \not\subseteq Q_k(U(\hat{D}))$ implies $\bigvee_i Q_i(D \cup \hat{D}) \not\subseteq V_i$, for every D, where $\hat{D}$ is the canonical database.* ∎

Theorem 4.1 reduces self-maintainability to the problem of deciding *DIFF* $\subseteq$ *INCON*, where *DIFF* is a query that represents the condition $Q_k(U(D \cup \hat{D})) \not\subseteq Q_k(U(\hat{D}))$, and *INCON* represents $\bigvee_i Q_i(D \cup \hat{D}) \not\subseteq V_i$. The rules that define these queries are listed in Table 1. Note that while negation is still used in *DIFF* and *INCON*, it only applies to constant predicates. The rules shown in Table 1 relate to Theorem 4.1 as follows: rules (A_{ij}) compute $\hat{D}$; (K_j) defines predicate r''_j for relation R_j in $D \cup \hat{D}$; (B'_i) represent the fact that $Q_i(D \cup \hat{D}) \not\subseteq V_i$; (D'_j) defines predicate r'_j for relation R_j in $U(D \cup \hat{D})$; (F_j) defines predicate $\hat{r}'_j$ for relation R_j in $U(\hat{D})$; (H_k) defines predicate $\hat{v}'_k$ for $Q_k(U(\hat{D}))$, the new state of view V_k that derives from $\hat{D}$ after the update; (I_k) expresses the fact that $Q_k(U(D \cup \hat{D})) \neq Q_k(U(\hat{D}))$. Based on the reduction from Theorem 4.1, the following algorithm generates a query that tests self-maintainability of view V_k.

Algorithm 4.1 *Generate Self-Maintainability Test*

Input: $Q_1, \ldots, Q_m$, where each Q_i is a CQ with no projection that defines v_i using $r_1, \ldots, r_n$ as input.

Output: A query that decides whether V_k is self-maintainable under U, using predicates $v_1, \ldots, v_m$ and $\delta r_1^-, \delta r_1^+, \ldots, \delta r_n^-, \delta r_n^+$ as input.

Method:

1. Generate rules for queries *DIFF* and *INCON* as shown in Table 1. Both queries use the 0-ary predicate *panic* for their query predicate.
2. Generate a query *TEST* that decides whether *DIFF* $\subseteq$ *INCON*. Return *TEST*.

∎

Note that instead of generating a query test as Step 2 of Algorithm 4.1 indicates, we could have solved *DIFF* $\subseteq$ *INCON* directly by using known algorithms in the literature ([G*94, Klu88]) for deciding containment of unions of CQ's with arithmetic comparisons. Even though these algorithms are more efficient than those for deciding containment of CQ's with negation, a naive way of applying them would require eliminating all constant EDB predicates (as Section 2 shows how). Unfortunately, the resulting complexity would still be exponential in the size of the views, because the expanded queries have exponential size. To show that *DIFF* $\subseteq$ *INCON* (and thus self-maintainability) can be decided in polynomial time, we show the existence of a nonrecursive query that can test *DIFF* $\subseteq$ *INCON* completely. In [Hu97], we show that this approach is possible precisely because negation used in *DIFF* and *INCON* only applies to constant predicates. Here are the key steps:

1. Translate *DIFF* $\subseteq$ *INCON* to a logical expression that involves the constant predicates used in the queries, rather than their extension.
2. Rewrite this (generally unsafe) expression to an equivalent safe expression which can be easily translated to a query (*TEST* in Algorithm 4.1) in safe, nonrecursive Datalog with negation and $\neq$ comparisons, or alternatively in SQL .

The supporting theorems and the algorithms that implement these steps are not shown here due to space limitation but can be found in [Hu97].

Using Algorithm 4.1, we can generate, at view-definition time, queries that test view self-maintainability at runtime. The time to generate these query tests does not depend on the instance of the views and update, but is generally exponential in the size of the view definitions. This complexity is not surprising, in light of the NP-completeness of checking query containment [CM77]. Evaluating the query tests obviously takes time polynomial in the size of the view instance and base update. However, it is important to optimize these tests further using compile-time query optimization techniques.

Example 4.2 Consider the definition of views V_1 and V_2 from Example 3.1 and consider the problem of testing self-maintainability of V_1 under the insertion of $r(a, b)$. Algorithm 4.1 generates a test which simplifies to the following query (using the 0-ary predicate *maintainable* as the query predicate):

$$\begin{aligned} p(Z) &\text{ :- } v_1(X,Y,Z) \\ q(Z) &\text{ :- } v_2(Y,Z) \;\&\; v_1(X,Y,Z') \\ \textit{depend} &\text{ :- } v_2(b,Z) \;\&\; \neg p(Z) \;\&\; \neg q(Z) \\ \textit{maintainable} &\text{ :- } \neg \textit{depend} \end{aligned}$$

or equivalently in SQL:

```
NOT EXISTS
(SELECT * FROM V2 WHERE V2.Y = b
 AND NOT EXISTS (SELECT * FROM V1
                 WHERE V1.Z = V2.Z)
 AND NOT EXISTS (SELECT * FROM V2 as V3, V1
                 WHERE V3.Z = V2.Z
                 AND V2'.Y = V1.Y))
```

∎

Table 1: Rules generated for the queries to compare in the reduction from Theorem 4.1.

Rules	Range	*DIFF*	*INCON*
(A_{ij})	$i = 1, \ldots, m$, $j = 1, \ldots, n_i$	$\hat{G}_{ij}$:- H_i	$\hat{G}_{ij}$:- H_i
(K_j)	$j = 1, \ldots, n$	r''_j :- r_j, r''_j :- $\hat{r}_j$	r''_j :- r_j, r''_j :- $\hat{r}_j$
(B'_i)	$i = 1, \ldots, m$		*panic* :- G''_{i1} & … & G''_{in_i} & $\neg H_i$
(D'_j)	$j = 1, \ldots, n$	r'_j :- r''_j & $\neg\delta r^-_j$, r'_j :- δr^+_j	
(F_j)	$j = 1, \ldots, n$	$\hat{r}'_j$:- $\hat{r}_j$ & $\neg\delta r^-_j$, $\hat{r}'_j$:- δr^+_j	
(H_k)		$\hat{H}'_k$:- $\hat{G}'_{k1}$ & … & $\hat{G}'_{kn_k}$	
(I_k)		*panic* :- G'_{k1} & … & G'_{kn_k} & $\neg\hat{H}'_k$	

Notation: H_i :- G_{i1} & … & G_{in_i} is the rule that defines view V_i; $\hat{G}_{ij}$ is the subgoal G_{ij} whose predicate r_l is replaced by predicate $\hat{r}_l$; G''_{ij} is G_{ij} whose r_l is replaced by r''_l; $\hat{H}'_k$ is H_k whose v_k is replaced by $\hat{v}'_k$; $\hat{G}'_{kj}$ is G_{kj} whose r_l is replaced by $\hat{r}'_l$; G'_{kj} is G_{kj} whose r_l is replaced by r'_l.

5 Maintaining Other Classes of Views

The techniques developed in the previous sections for maintaining a special class of views (defined by conjunctive queries without projection) have, in fact, much wider applicability. In this section, we show how to extend them to conjunctive-query views with projection and partial copies. Other extensions are possible (e.g., for queries with arithmetic comparisons, unions of conjunctive queries, and queries over base relations constrained by dependencies) but are not described here due to space limitation.

5.1 Views with Projections

Consider views defined by conjunctive queries where some variables used in a rule's body do not appear in the rule's head. A view where some attributes have been projected out looses information, and from an instance of the view, there is no unique way of "reconstructing" a minimal database. The notion of canonical database from Section 3 must be revised to capture this nonuniqueness. So, we redefine our new canonical database $\hat{D}$ as follows: for each V_i, since each tuple in V_i that matches Q_i's head provides a substitution for only some of the variables in Q_i's body, this substitution is extended to the remaining variables by binding each of them to a *new symbol*; the ground atoms obtained after making this extended substitution into Q_i's body are included in $\hat{D}$.

A tuple in $\hat{D}$ that contains a new symbol represents a fact involving some object whose value is not known. This value could be any of the known constants from the instance of the views or the update instance, or could be some constant not in any of those instances. Thus, if we consider all the symbol mappings h that map each of the new symbols to either one of themselves or a known constant, then $\hat{D}$ represents not a single database but a class of possible databases, each of which is obtained by applying some substitution h to $\hat{D}$. We can show that there is always a mapping h such that database $h(\hat{D})$ is consistent with all the views. Such a mapping is said to be *consistent.* The following example illustrates the nonuniqueness of minimal databases due to projections in views.

Example 5.1 Consider the view definition $v(X,Z)$:- $s(X,Y)$ & $s(Y,Z)$ where Y has been projected out. Consider the instance $V = \{(d,c)\}$ and the insertion of (a,b) to S. The canonical database $\hat{D}$, obtained as $S = \{(d,y),(y,c)\}$ where y is a new symbol, actually can be interpreted in five possible ways (by mapping y to either y, a, b, d, or c): $S = \{(d,y),(y,c)\}$, $S = \{(d,a),(a,c)\}$, $S = \{(d,b),(b,c)\}$, $S = \{(d,d),(d,c)\}$, or $S = \{(d,c),(c,c)\}$. The last two databases are not consistent with V, since they respectively derive tuples (d,d) and (c,c), which are not in the view. Among the remaining consistent databases, after the insertion, the second one derives tuple (d,b) not derived by the first one. Thus, view V is not self-maintainable under the insertion of (a,b) to S. ▮

Once a consistent database $h(\hat{D})$ is found, we can apply the same idea as in Section 3 to maintain a view if the view is self-maintainable: propagate an update to the view using $h(\hat{D})$ as the actual database. Similarly, the self-maintainability question is settled by extending the reduction from Section 4 to take into account the nonuniqueness of a minimal consistent database. We can now informally state a theorem analogous to Theorem 4.1: V_k is self-maintainable under U if and only if (1) the new state of V_k does not depend on which consistent database $h(\hat{D})$ we use as the actual database, and (2) for each consistent $h(\hat{D})$, every database that contains $h(\hat{D})$ and that is consistent with all the views must derive the same state for V_k after the update, as $h(\hat{D})$.

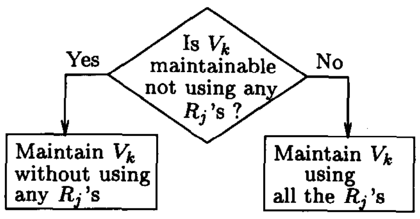
To sum it up, we obtain algorithms similar to Algorithms 3.1 and 4.1 except that they use a minimal database $h(\hat{D})$ that is consistent with all the views, instead of just $\hat{D}$. While the use of projection in views seems to make the problem considerably harder since the number of consistent mappings h can be exponential in the worst case, results from [Hu96] suggest that it does not have to be so. For example, [Hu96] showed that even with projection, self-maintainability of a single conjunctive-query view with no self-join can be efficiently decided with a simple query. Thus, an important future direction is to further refine our techniques and identify restrictions on the view that allow the problem to be solved efficiently.

5.2 Partial Copies

Suppose a log of the most recent updates on a base relation R is kept at the warehouse. Unlike a full copy of R stored at the warehouse, the most the log can tell us about R is that R must include certain tuples (say represented by set R^+) but exclude others (say R^-). We call this information about relation R a *partial copy*. Thus, given R^+ and R^-, the question is how to take full advantage of the additional information in view self-maintenance. A solution can be obtained by simply revising both Definition 3.1 of the canonical database to additionally include R^+ and Theorem 4.1 to use $R \cap R^- \neq \emptyset$ as another inconsistency condition on the right hand of the implication. Algorithms similar to Algorithms 3.1 and 4.1 can be obtained the obvious way. Furthermore, the solutions can be extended in a straightforward manner to logs that contain the most recent updates on more than one base relation. View self-maintenance with partial copies has the same complexity as with regular materialized views.

6 Maintaining Views with Partial Access to the Base Relations

As stated, the main motivation behind self-maintenance is in maintaining a warehouse efficiently by minimizing the use of base relations. Thus, in strict view self-maintenance, we attempt to maintain the views using information that can be obtained strictly locally from the warehouse, namely the materialized views and the update. When a view is not self-maintainable in the strict sense, an obvious strategy is to fall back to the "normal" but expensive view maintenance mode with unrestricted access to the base relations, as depicted in Figure 2(a). However, instead of switching to the normal maintenance mode immediately, we may be able to use some (but not necessarily all) of the base relations to successfully maintain the view. In fact, there are many cases where a view is not self-maintainable in the strict sense but can be maintained using some of the base relations. Thus, a more refined strategy based on generalized self-maintenance can be used instead, as illustrated in Figure 2(b). Note that in this strategy, the choice of which subset base relations to use next is left open. How to make the optimal choice is an important area for future research.


Is V_k maintainable not using any R_j's ?
Yes
No
Maintain V_k without using any R_j's
Maintain V_k using all the R_j's
(a) Under strict self-maintenance

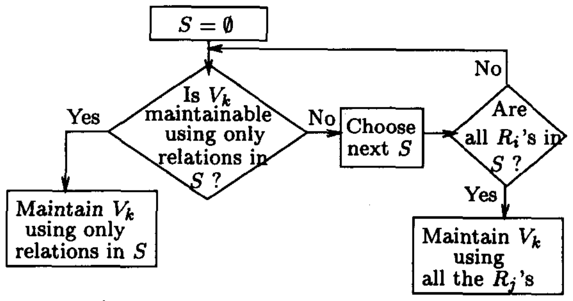

$S = \emptyset$
Is V_k maintainable using only relations in S ?
Yes
No
Choose next S
Are all R_i's in S ?
No
Yes
Maintain V_k using only relations in S
Maintain V_k using all the R_j's
(b) Under generalized self-maintenance

Figure 2: Strategies for Efficient Maintenance.

In the following, we show how to solve the generalized self-maintenance problem. There is a close resemblance between ***allowing access to a base relation*** and ***having a copy of the base relation materialized at the warehouse***. In fact, if we assume that:

- The materialized views are ***simultaneously*** updated, that is, the required updates to each view are determined prior to updating any view, and
- The base relations are accessed in a state that reflects update U but no other later updates (assuming that the warehouse received updates in the order they are applied to the database),

then, the generalized self-maintenance problem can be treated as a strict self-maintenance problem where ***a copy of the given base relations is available at the warehouse***, with the exception that the actual base relations and the copy only differ by the update.

Example 6.1 Consider a warehouse with a single view V_1 defined as in Example 3.1. Assume we can access base relation S but not R or T. Consider an update with δR^-, δR^+, δS^-, δS^+, δT^-, and δT^+. The maintenance expression and maintainability test for this generalized self-maintenance problem can be obtained as follows. Consider the strict self-maintenance problem with both view V_1 and a view V_2 that is a copy of S. The solutions to this problem use predicates v_1 and v_2 as input. Replace every occurrence of v_2 with a new predicate s' defined by the following rules:

$$s'(Y,Z) \text{ :- } s(Y,Z) \ \& \ \neg\delta s^+(Y,Z)$$
$$s'(Y,Z) \text{ :- } \delta s^-(Y,Z)$$

Predicate s' represents the state of relation S prior to the given update. ▮

Thus, results for the strict self-maintenance problem can be carried over by simply replacing every reference to the "copy" of a base relation by a reference to its "before image". In practice, allowing access to a base relation when maintaining a materialized view must be handled carefully. When a base relation is asynchronously updated by the source, it may be read by the warehouse in a different state than what is assumed by the warehouse. This situation may lead to erroneous updates to the warehouse, as reported in [Z*95]. Thus, a warehouse system that uses generalized self-maintenance must either allow access only to base relations that change in lock step with the warehouse, or combine our techniques with the compensation techniques developed in [Z*95].

7 Conclusion and Future Work

We have given algorithms that test view maintainability and incrementally maintain a view in response to a base update, based on the current state of all the views in the warehouse and of a specified subset of the base relations. We improve significantly on previous work, because our methods allow us to take full advantage of all the views stored at the warehouse and to handle base updates that consist of arbitrary mix of insertions and deletions. The techniques used in obtaining the algorithms are applicable to a wide variety of views, and in some cases, allow us to generate tests and maintenance expressions in the form of SQL queries. We intend to use these algorithms as the basis for maintaining a warehouse efficiently. The practicality of our approach can be enhanced with a better ability to optimize the queries generated by the algorithms, a better strategy for selecting which subset of the base relations to access if a view turns out to be not maintainable, and by extending our techniques to deal with multi-set semantics.

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