Query Decomposition and View Maintenance for Query Languages for Unstructured Data

Dan Suciu

AT&T Research, USA
suciu@research.att.com

Abstract

Recently, several query languages have been proposed for querying information sources whose data is not constrained by a schema, or whose schema is unknown. Examples include: LOREL (for querying data combined from several heterogeneous sources), W3QS (for querying the World Wide Web); and UnQL (for querying unstructured data).

The natural data model for such languages is that of a rooted, labeled graph. Their main novelty is the ability to express queries which traverse arbitrarily long paths in the graph, typically described by a regular expression. Such queries however may prove difficult to evaluate in the case when the data is distributed on several sites, with many edges going between sites. A typical case is that of a collection of WWW sites, with links pointing freely from one site to another (even forming cycles). A naive query shipping strategy may force the query to migrate back and forth between the various sites, leading to poor performance (or even non-termination). We present a technique for *query decomposition*, under which the query is shipped exactly once to every site, computed locally, then the local results are shipped to the client, and assembled here into the final result. This technique is efficient, in that (a) only data which is part of the final result is shipped from the data sites to the client site, and (b) the total work done locally at all sites does not exceed that needed for computing the (unoptimized) query on a centralized version of the database.

We also show that the query decomposition technique can be adapted to derive a simple view maintenance method, for two forms of updates which we introduce for the graph data model.

**Proceedings of the 22nd VLDB Conference
Mumbai(Bombay), India, 1996**

1 Introduction

Several database query languages have been proposed recently, in which the data model is that of a labeled tree, or, more generally, a labeled, rooted graph. Lorel [QRS+95] was designed in conjunction with the Tsimmis project [PGMW95], to query data from information sources which do not impose a rigid structure, or whose structure is not completely known. UnQL [BDS95, BDHS96a], was motivated by the need to query self-describing databases, like the ACeDB databases popular among biologists, and also for data sources with unknown structure. Authors of both languages suggest their suitability to query information sources on the World-Wide Web. W3QS [KS95] is a query language specifically designed for querying the World-Wide Web. It was motivated by the fact that today's web search engines are restricted to *content based* queries, which select a single page based on its content: W3QS was designed to ask *structure based* queries, addressing the hypertext organization itself. Another system developed for querying hypertext structures is presented in [BK90].

The main common feature of these languages is their ability to follow arbitrarily long sequences of links. We illustrate this feature by giving a typical query, which we will use throughout this paper, based on an WWW example.

Example 1.1 Suppose that we want to fetch all publications from the Computer Science Department at the University of California at San Diego, whose home page is `http://www.ucsd.edu`. The html structure we need to search is illustrated in Figure 1. It is convenient to view html data as a rooted, directed graph: nodes correspond to pages, edges to links, and each page is a *set* of all its outgoing links[1]. The prob-

[1]In Figure 1, and throughout the paper, we will ignore the node content: thus our trees will have labeled edges, not nodes. In the case of html data we justify our choice by the fact that we focus on structural, rather than content based queries. See also [BDHS96a] for more justification of this model.

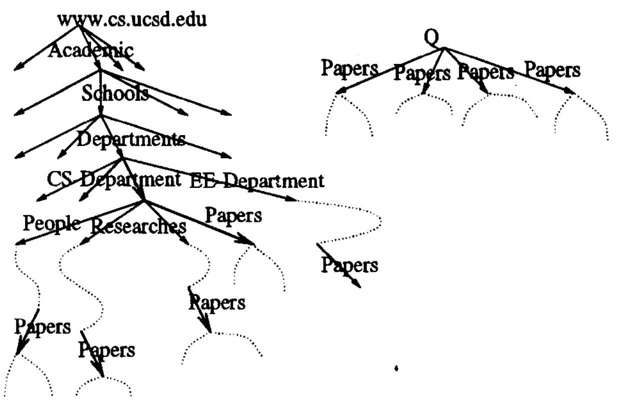


Figure 1: A fragment of `http://www.ucsd.edu`, and the result of a query Q.

lem in expressing this query in a relational or object-oriented query language is that we don't know in advance how many links to follow from the root to the *CS-Department* link, and from here to the *Papers* link. In a slightly modified UnQL syntax this query is expressed as:

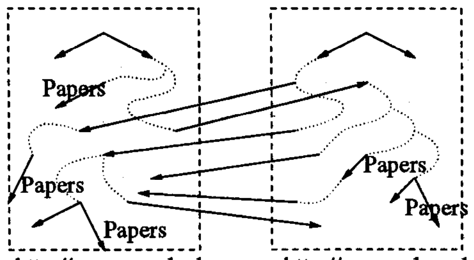
select *"Papers".t*
where $\_*$. *"CS-Department"*.$\_*$. *"Papers".t* in DB

Here DB stands for the "database", in this example the `http://www.ucsd.edu` site. The symbol _ denotes any label, while $*$ placed after some label means that it can be repeated 0 or more times. Thus $\_*$ means any sequence of labels. In short, the query starts from the root, follows all paths to an edge labeled *"CS Department"*, from there follows all paths to an edge labeled *"Papers"*, and returns the set of all subtrees thus found. Of course, the same query can be applied to other university sites, to retrieve papers published in their Computer Science Departments.

This example illustrates the main feature separating query languages like UnQL, LOREL, W3QS from relational or object-oriented query languages: their ability to follow arbitrarily long sequences of links. Typically such sequences are described by regular expressions. In particular we can compute the transitive closure of the link relation. Following [BDHS96a] we will call such languages ***unstructured query languages***. Their natural data model is that of a rooted, directed graph, with labels attached to vertices, or edges, or both. Traditional relational databases are captured by this model as the special case of trees of a fixed depth ([BDHS96a], see also Section 3 in this paper). Over relational databases however, the expressive power of such languages is no more powerful than that of the relational algebra [BDHS96a].

In this paper we discuss two problems in connection with such languages: query decomposition and view maintenance. We describe these two problems next.

Query decomposition Information sources sometimes reside on a number of different sites. The database fragment in any one site may have a large number of links leading to other sites. To illustrate,



http://www.ucsd.edu http://www.sdsc.edu

Figure 2: The information source is distributed on two sites.

suppose we want to consider UCSD's web site in Example 1.1, in conjunction with the web site of the San Diego Supercomputer Center, `http://www.sdsc.edu` (Figure 2). A typical query Q may start at the UCSD web site, but then will traverse back and forth links between the two sites, as it traverses sequences of links. E.g. the *"CS-Department"* may have links to projects at the SDSC site, while from here we may follow links back to the UCSD site, etc: at any point, we may encounter a *"Papers"* link. If query shipping is used for query evaluation, this implies that the query has to migrate between the two sites an arbitrary large number of times. The ***query decomposition*** problem requires the query Q to be decomposed into two queries which can be computed independently on the fragments DB_1 and DB_2 of the database residing at the two sites, and their results combined at the client's site to yield $Q(DB)$. Of course, this problem can be solved best when additional knowledge about the data distribution is available, e.g. that the *"CS-Department"* link is unique and that it resides at the UCSD site. But our purpose here is to see how much we can accomplish without any knowledge about the data distribution: this is useful, e.g. when browsing an unfamiliar information source. We will describe a technique which translates any query Q in a certain fragment of an unstructured query language, into a query Q' which is ***decomposable***, meaning that $Q'(DB)$ can be computed by computing $Q'(DB_1)$ and $Q'(DB_2)$ independently, then gluing the two results at the client's site. Moreover, this decomposition is efficient: computing $Q'(DB_1)$ and $Q'(DB_2)$ is no more expensive than computing $Q(DB)$, using some naive evaluation strategy, when DB is centralized. However an ***optimized*** evaluation of $Q(DB)$ could be more efficient on a centralized version of DB than that of computing $Q'(DB_1)$ and $Q'(DB_2)$ separately.

View maintenance A view V is just the result of some query Q applied to the database: $V = Q(DB)$. The view is ***materialized*** when the result V is stored at the client site for future use. The view maintenance problem consists in finding efficient techniques for updating the view V at the client site when the database is incrementally updated, say $DB' := DB \mathbin{+\!\!+} \Delta$, where Δ is much "smaller" than DB ($\mathbin{+\!\!+}$ defined in Section 3; for the case of relational databases represented as trees, insertion of an element into a set can be modeled as a special case of $\mathbin{+\!\!+}$). Ideally one would like the view update to depend only on the old view and

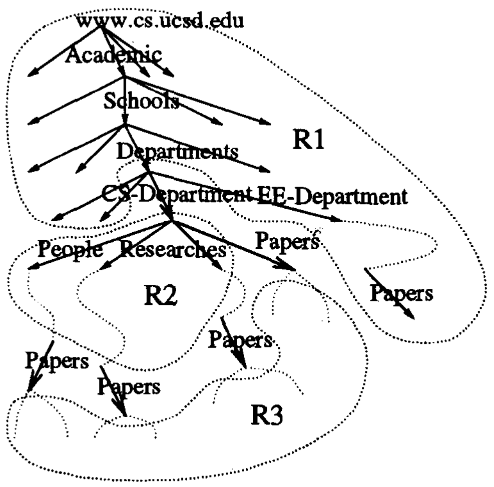


Figure 3: The query Q logically partitions the database in three regions, $R1, R2, R3$.

the increment, $V' := f(V, \Delta)$, not on DB. View maintenance techniques for relational databases have been extensively discussed, see e.g. [GL95] for a list of references. In the case of unstructured data however, we face an additional problem: namely that of informing the client site *where* the update has taken place. The "where" information is taken for granted in the case of a relational query language. E.g. consider a relational database with three relations R_1, R_2, R_3, and the view $V = R_1 \cup \sigma_{1=3}(R_2)$. In classic view maintenance algorithms the view will be updated differently, depending on whether R_1, R_2 or R_3 has been updated. E.g. when $R_2' := R_2' \cup \Delta$, then the view becomes $V' := V \cup \sigma_{1=3}(\Delta)$, but when R_3 gets updated the view does not need to be changed. In an "unstructured" database there is no static partition of the data into distinct relations, but the partition is done dynamically, when the view is computed. The query in Example 1.1 defines a view which logically partitions the database into three regions, as in Figure 3. The regular expression occurring in this query has an equivalent automaton with three states: the three regions correspond precisely to these states. View maintenance will be done differently after an update in region $R1$ than after updates in regions $R2$ or $R3$ respectively. Of course a view defined by another query may partition DB in a different way. In this paper we develop algebraic techniques which transform any query Q in a certain fragment of an unstructured query language into a query Q' (the same as in the query decomposition problem), such that (1) $Q(DB)$ can be computed easily from $Q'(DB)$, (2) the computation of $Q'(DB)$ is no more expensive than a naive computation of $Q(DB)$, and (3) $Q'(\Delta)$ "encapsulates" the information about which region has been affected by the update, and, furthermore: $Q'(DB +\!\!+ \Delta) = Q'(DB) +\!\!+ Q'(\Delta)$.

1.1 Relation to previous work

We choose to describe our techniques in the context of the query language UnQL [BDHS96a, BDHS96b]. The language relies on a data model of rooted, labeled graphs, which was first described in [BDS95]. The same paper describes bisimulation between rooted, labeled graphs, showing how this data model subsumes both the relational and the nested relational one. Finally it introduces the basic algebraic operations associated to this data model, including tree concatenation (denoted $+\!\!+$ in this paper), and *vext* ("vertical" *ext*), observing that *vext* works nicely in conjunction with cyclic structures. The language UnQL is introduced in [BDHS96a, BDHS96b], as a declarative language with pattern matching, and is shown to be equivalent to an algebraic language UnCAL, centered around $+\!\!+$ and a generalized version of *vext* called *gext* ("generalized" *ext*). [BDHS96a, BDHS96b] prove a number of algebraic laws for *gext* which are intended to be used for query optimizations. In this paper we use these laws and additional ones, dealing with cyclic data, in order to derive our query decomposition techniques. We also need to add some additional operations to UnCAL: namely we consider labeled graphs with several roots as first class objects in the language (much like the n-trees in [Cou83, pp 134]), and extend UnCAL with ***graph juxtaposition***, see Section 4. The equations associated to these new operations, much in the spirit of [Cou83, pp 135], are simpler however than those for $+\!\!+$ and *gext*.

1.2 Applicability, assumptions and limitations

Query language Although we describe our techniques in the context of the query language UnQL [BDS95, BDHS96a], they apply equally well for decomposing queries in (certain fragments of) LOREL or W3QS. It is important to notice however that during both query decomposition and view maintenance, we use in a critical way some of the particular language constructs in UnQL, such as tree concatenation, $t +\!\!+ t'$ (denoted @ in [BDHS96a]). Moreover we note that our techniques apply to some UnQL queries which are not expressible in LOREL or W3QS.

Restrictions on the query language Our query decomposition technique works only for UnQL queries satisfying two restrictions. The first requires the query to be "monotone". In particular we cannot handle set difference, or incremental deletions. This is an expected limitation, because our data model has a set semantics: in particular flat trees are just sets. For the case of relational databases, it is known [GL95] that bag semantics rather than set semantics is needed in order to do algebraic view maintenance for deletions. The second restriction requires the query to be join-free. This limitation is far less severe than it sounds, because in unstructured query languages the focus is on queries traversing long link sequences, which are join-free: as in object-oriented databases, most joins are replaced by link traversals. In fact *all* queries in W3QS [KS95] are join-free, while most interesting LOREL and UnQL examples [QRS+95, BDHS96a] are join-free too.

No knowledge about data distribution Our query decomposition method works without assuming any knowledge about how the data is distributed on the sites. While sometimes this may be useful, we consider it to be the most serious limitation of our approach. In most cases minor knowledge about the data sources could help one decompose a query much better than we currently do. As part of future work, we plan to investigate how knowledge about data sites can be incorporated into our query decomposition method.

Updates Our view maintenance techniques work for two kinds of updates: *insertions* and *replacements*. An insertion, in notation $DB +\!\!+ \Delta$, means that some new subtree Δ is inserted at some particular node v in DB (Section 3): the root of Δ will be "merged" with v. This is a monotone operation, in that the resulting database has at least as many edges as the original one. The second kind of update, *replace*, allows us to replace a subtree rooted at a given node with another tree. This is not a monotone operation and can, to some extent, model deletions, like those in a relational databases.

Update notification Our view maintenance technique requires that the server informs the client whenever a page (or, in general, a node in the graph) is updated. We refer the reader to [BD96, DB96] for update notification techniques for the WWW.

Site selection and query capabilities We do not address the problem of site selection. Instead, we always assume that the database is stored at a fixed, known, relatively small number of sites, say $s_1, \dots, s_k$, which may have links between them. This assumption could be easily enforced, in the case of web sites, by simply ignoring all links pointing outside the set $s_1, \dots, s_k$. Also, we assume no knowledge about the semantics of data stored at these sites, or about their query capabilities: in reality some sites may offer only restricted access, e.g. using a keyword search. For the case of conjunctive queries over relational databases, [LMSS96, LRU96] discuss techniques which can efficiently select the relevant sites and also use the limited query capabilities of such sites.

2 An Example

We illustrate the problems, and the techniques we propose to solve them, for the query in Example 1.1. It is easier to describe first the view maintenance technique, so we will start with that.

For view maintenance, we require a database DB to have all its updatable nodes explicitly marked. In an extreme case, all nodes could be marked as updatable, but adopting this approach could lead to an increased storage requirement for some materialized views. When the view $V = Q(DB)$ is first computed, the semantics of the UnQL query Q is such that the result V encapsulates some (or all) markers of the updatable pages in DB. More, each such marker is tagged with the name of the (dynamically computed) region where it was encountered. Suppose now that the database DB is updated, say at a page marked X, in that a link to a new subgraph Δ is added to that page: in notation $DB' := DB +\!\!+_X \Delta$. The server notifies the client about the update, by sending X and Δ. The client "looks up" the marker X in its view, and, if present, reads the tag of the region where it occurred (R_1, R_2, or R_3), then updates the view dynamically, as shown next. In our particular example, in all three cases the view needs to be updated to $V' := V \cup \Delta'$, where Δ' is as follows:

1. If Δ is inserted before any *"CS-Department"* edge (in region $R1$ of Figure 3), then $\Delta' = Q_1(\Delta)$, where $Q_1(\Delta)$ is:

   select *"Papers".t*
   where _ * . *"CS-Department"*._ * . *"Papers".t*
   in Δ

2. If Δ is inserted between a *"CS-Department"* edge and a *"Papers"* edge (region $R2$ of Figure 3), then $\Delta' = Q_2(\Delta)$, where $Q_2(\Delta)$ is:

   select *t*
   where _ * . *"Papers".t* in Δ

3. Finally, if Δ is inserted after a *"CS-Department"* and after a *"Papers"* edge, then[2] $\Delta' = Q_3(\Delta) = \Delta$.

While our informal description here sounds rather procedural, the technique, as described further in this paper, is fully algebraic.

We illustrate next query decomposition. Assume that the database DB is distributed on two different sites s and s', with arbitrary many links between them: then a path of the form

_ * . *"CS-Department"*._ * . *"Papers"*

may travel back and forth between s and s' an arbitrary number of times. Since we assume no a priori knowledge about how the data is structured and/or partitioned on the two sites, we should be prepared to deal with more than one *"CS-Department"* links, with any number of *"Papers"* links, and with any possible distribution of these links on the two sites. To decompose the query we need to have a list of all input and output links at each site (Figure 4 (a)). Intuitively, our method starts by decomposing the query Q into the three queries Q_1, Q_2, Q_3 described above, and then applies each of these three queries to every entry point at each site; this is illustrated by the six inputs $Q1(X), \dots, Q3(Y)$ at the site s' and the six inputs $Q1(U), \dots, Q3(V)$ at the site s in Figure 4 (b). Now we observe that each of these queries logically partitions the database into up to three regions (Q_2 defines only two, while Q_3 only one). We want to tag each output link with the corresponding region name where it was found. But note that the three partitions are independent, and each output link may be "found" in more than one region, say in R_3 by Q_1, and in R_2 by Q_2. Hence we replicate each output link 3 times, once for each region where it may be found. This is illustrated by the six outputs at sites s and s'

[2]In this case the correct view update is of the form $V' := V +\!\!+ \Delta'$; see Section 8.

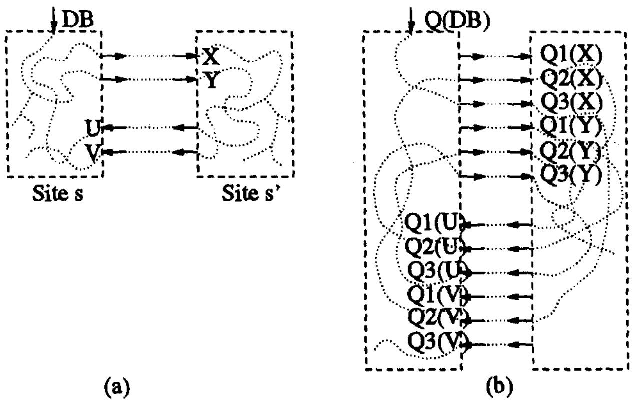


Figure 4: Query decomposition for two sites.

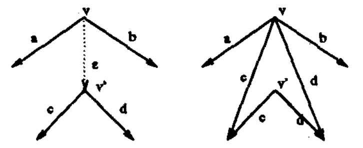


Figure 5: Intuitive meaning of an ε edge.

in Figure 4 (b). In the next step the results of the two independent computations are shipped to the client, where they are combined, by gluing the outputs with their "corresponding" inputs.

3 Data Model

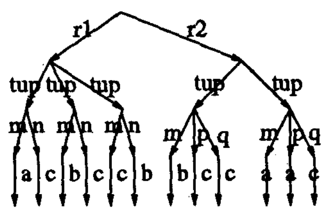
As mentioned earlier, we adopt the data model in [BDS95, BDHS96a]. Unlike the Tsimmis data model [PGMW95, QRS+95] it does not have object id's, and has two additional features, markers and ϵ-edges, which we use in a critical way for both query decomposition and view maintenance. We briefly review here the main ideas and refer the reader to [BDHS96b] for a detailed description of the data model.

Rooted graphs Let *Label* be the universe of all labels: it includes all strings, numbers, booleans, etc. A database is modeled as a **rooted graph** (i.e. a graph with a distinguished node called the *root*), whose edges are labeled with elements from $Label \cup \{\varepsilon\}$. Here ε is a special label, denoting an "empty" symbol: whenever two vertices v, v' are connected by an ϵ edge, the intended meaning is that all edges emerging from v' should also emerge from v, see Figure 5.

Trees Trees form a particularly interesting subset of the rooted graphs, and they suffice to represent sets and records. Figure 6 contains an example of a relational database and its representation as a tree. The following is a syntax for trees:

$$Tree \ ::= \ \{\} \mid \{Label \Rightarrow Tree\} \mid Tree \cup Tree$$

Abbreviating $\{a_1 \Rightarrow t_1, \ldots, a_n \Rightarrow t_n\}$ for $\{a_1 \Rightarrow t_1\} \cup \ldots \cup \{a_n \Rightarrow t_n\}$, and $\{a\}$ for $\{a \Rightarrow \{\}\}$, the example in Figure 6 is written as:



r1:

m	n
a	c
b	c
c	b

r2:

m	p	q
b	c	c
a	a	c

Figure 6: A relational database represented as a tree.

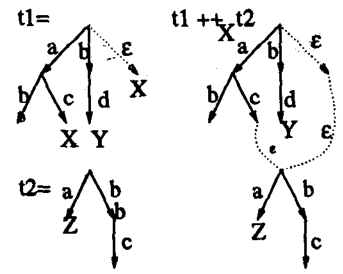


Figure 7: Illustration of $t_1 +\!\!+_X t_2$.

$$\begin{aligned}\{r1 \Rightarrow \{&tup \Rightarrow \{m \Rightarrow \{a\}, n \Rightarrow \{c\}\},\\ &tup \Rightarrow \{m \Rightarrow \{b\}, n \Rightarrow \{c\}\},\\ &tup \Rightarrow \{m \Rightarrow \{c\}, n \Rightarrow \{b\}\}\},\\ r2 \Rightarrow \{&tup \Rightarrow \{m \Rightarrow \{b\}, p \Rightarrow \{c\}, q \Rightarrow \{c\}\},\\ &tup \Rightarrow \{m \Rightarrow \{a\}, p \Rightarrow \{a\}, q \Rightarrow \{c\}\}\}\}\end{aligned}$$

We emphasize that this data model has set semantics. E.g. the trees $\{a, b \Rightarrow \{c\}, b \Rightarrow \{c\}\}$ and $\{a, b \Rightarrow \{c\}\}$ are considered equal.

Markers In addition to the edge labels, some of the leaves of a graph are allowed to be labeled with special symbols, denoted $X, Y, \ldots$, called *markers*. Unlike labels, markers are not part of the information content of the database, but are used to control (1) where updates take place, and (2) how to connect fragments of a distributed database. They share some similarities with the object id's used in the Tsimmis data model [PGMW95, QRS+95]. Markers allow us to define the concatenation operation $+\!\!+_X$: given two graphs t_1, t_2 and a marker X, $t_1 +\!\!+_X t_2$ denotes the database obtained by drawing ε edges from all leaves labeled X in t_1 to the root of t_2. All occurrences of the old marker X in t_1 disappear in $t_1 +\!\!+_X t_2$. But all other markers in t_1 remain in $t_1 +\!\!+_X t_2$, as well as all markers from t_2, see Figure 7.

Graphs with m inputs, n outputs To capture the connection between fragments of a distributed database, we generalize from single input (the root of the tree) to m inputs. Also we call the markers on the leaves the *outputs* of the tree, thus reaching the notion of a tree (graph) with m inputs and n outputs. Formally, let $\mathcal{X} = \{X_1, \ldots, X_m\}$ and $\mathcal{Y} = \{Y_1, \ldots, Y_n\}$ be two finite sets of markers, $m \geq 0, n \geq 0$. A database

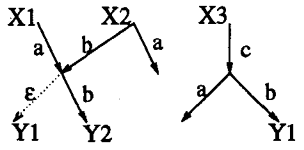


Figure 8: A database with inputs $\mathcal{X} = \{X_1, X_2, X_3\}$ and outputs $\mathcal{Y} = \{Y_1, Y_2\}$.

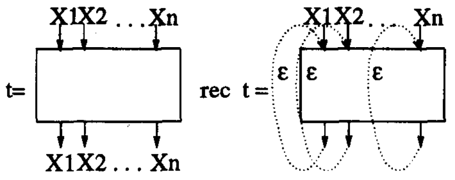


Figure 9: Illustration of rec t where t has inputs and outputs $\{X_1, \ldots, X_n\}$.

with inputs $\mathcal{X}$ and outputs $\mathcal{Y}$ is a graph with edges labeled with elements from *Label* $\cup \{\varepsilon\}$ and in which some leaves may be labeled with markers in $\mathcal{Y}$ (as before), and with m distinguished roots associated to $X_1, \ldots, X_m$. We extend the syntax for trees to that of trees with inputs $\mathcal{X}$ and outputs $\mathcal{Y}$, $Tree^{\mathcal{X}}_{\mathcal{Y}}$, and to trees with a single, anonymous input and with outputs $\mathcal{Y}$, $Tree_{\mathcal{Y}}$:

$$\begin{aligned} Tree^{\mathcal{X}}_{\mathcal{Y}} &::= (X_1 := Tree_{\mathcal{Y}}; \ldots; X_m := Tree_{\mathcal{Y}}) \\ Tree_{\mathcal{Y}} &::= \{\} \mid \{Label \Rightarrow Tree_{\mathcal{Y}}\} \mid Tree_{\mathcal{Y}} \cup Tree_{\mathcal{Y}} \\ &\quad \mid Y_j (j = 1, n) \end{aligned}$$

Figure 8 contains an example of a graph with inputs X_1, X_2, X_3 and outputs Y_1, Y_2, which is written as $(X_1 := \{a \Rightarrow Y_1 \cup \{b \Rightarrow Y_2\}\}; X_2 := \{b \Rightarrow \{Y_1 \cup \{b \Rightarrow Y_2\}\}, a\}; X_3 := \{c \Rightarrow \{a, b \Rightarrow Y_1\}\})$. By convention, there exists a single graph with inputs $\mathcal{X} = \emptyset$, namely the null graph, denoted () (no nodes, no edges: this is different from the empty set, {}, which has one node, no edges). We extend the ++ operation as follows. For $t_1 \in Tree^{\mathcal{X}}_{\mathcal{Y}}$ and $t_2 \in Tree^{\mathcal{Y}}_{\mathcal{Z}}$, $t_1 +\!\!+_{\mathcal{Y}} t_2$ is the tree in $Tree^{\mathcal{X}}_{\mathcal{Z}}$ obtained by drawing an ε edge from every output Y_i in t_1 to the input Y_i in t_2, $i = 1, n$.

Recursive definitions For $t \in Tree^{\mathcal{X}}_{\mathcal{X}}$, $\text{rec}_{\mathcal{X}}\ t$ denotes a graph obtained by adding an ε-edge from every output X_i to the input X_i, $i = 1, m$ (Figure 9). Intuitively $\text{rec}_{\mathcal{X}}\ t$ is $t +\!\!+_{\mathcal{X}} t +\!\!+_{\mathcal{X}} t +\!\!+_{\mathcal{X}} \ldots$ E.g. $\text{rec}_X\ (X := \{a \Rightarrow X\})$ defines a loop labeled a.

Equality The notion of equality on rooted graphs is that of **bisimilation** [BDHS96a, BDS95]. In a nutshell, two rooted graphs are bisimilar if, after (possible infinite) unfolding, ε-edge removal, and duplicate subtree elimination at each node, the two resulting trees are equal. When restricted to tree representations of sets and records, as in Figure 6, the bisimilarity relation is exactly the set and/or record equality.

ε-Edges ε-Edges are introduced for convenience of notation: most operations, like ++ or rec, are easier defined in terms of ε edges, than without them. On the other hand the reader may have noticed that, in some

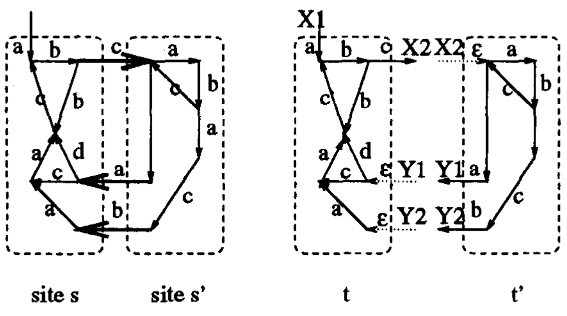


Figure 10: A distributed database.

sense, they are redundant: every rooted graph with ε edges can be shown to be bisimilar to a rooted graph without ε edges [BDS95, BDHS96a]. But this only works when the graphs have no markers. E.g. the tree $\{a\} \cup X$ represented with an ε edge as $\{a \Rightarrow \{\}, \varepsilon \Rightarrow X\}$, is not equivalent to any tree without ε edges.

Notation As in [BDHS96a], by abuse of notation we call "trees" all graphs, even when they have cycles. Thus $Tree_{\mathcal{Y}}$ denotes the set of all databases with a single anonymous input and outputs $\mathcal{Y}$, and $Tree^{\mathcal{X}}_{\mathcal{Y}}$ that of all databases with inputs $\mathcal{X}$, outputs $\mathcal{Y}$. When $\mathcal{Y} \subseteq \mathcal{Y}'$, then $Tree_{\mathcal{Y}} \subseteq Tree_{\mathcal{Y}'}$ and $Tree^{\mathcal{X}}_{\mathcal{Y}} \subseteq Tree^{\mathcal{X}}_{\mathcal{Y}'}$.

Canonical form Every database DB in $Tree_{\mathcal{Y}}$ can be expressed (not necessarily uniquely) as:

$$X_1 +\!\!+_{\mathcal{X}} \text{rec}_{\mathcal{X}}\ (X_1 := t_1; \ldots, X_m := t_m) \qquad (1)$$

for some set $\mathcal{X} = \{X_1, \ldots, X_m\}$, where each of t_i, $i = 1, m$ is cycle free.

4 Representing Distributed Databases

We will illustrate how a distributed database can be represented in our notation, using the example in Figure 10, where DB is stored on two sites s, s'. We start by cutting the cross links, and inserting markers in place: X_1, X_2, Y_1, Y_2, where X_1 is the input to the old root. Let t and t' be the two fragments of the database residing at the two sites: here $t \in Tree^{\mathcal{X}}_{\mathcal{Y}}$ and $t' \in Tree^{\mathcal{X}'}_{\mathcal{Y}}$, with $\mathcal{X} \stackrel{\text{def}}{=} \{X_1, Y_1, Y_2\}, \mathcal{X}' \stackrel{\text{def}}{=} \{X_2\}$, and $\mathcal{Y} \stackrel{\text{def}}{=} \mathcal{X} \cup \mathcal{X}' = \{X_1, X_2, Y_1, Y_2\}$. Here X_1 denotes the root of DB. Then $DB \in Tree$ is recaptured as:

$$X_1 +\!\!+_{\mathcal{Y}} (\text{rec}_{\mathcal{Y}}\ (t; t'))$$

where $(t; t')$ is the *juxtaposition* of t and t' (disjoint graph union). The $\text{rec}_{\mathcal{Y}} \ldots$ construct redraws the cross links, while $X_1 +\!\!+_{\mathcal{Y}} (\ldots)$ selects only the input labeled X_1 as the unique, anonymous input of the database.

In general, every database DB stored at, say, two sites s, s' can be expressed in a canonical form as $X_1 +\!\!+_{\mathcal{X}} \text{rec}_{\mathcal{X}}\ (t; t')$, with t stored at site s and t' at site s'. Here t and t' need not be cycle free.

5 Transaction Language

Our view maintenance technique works in conjunction with a simple transaction language in the spirit

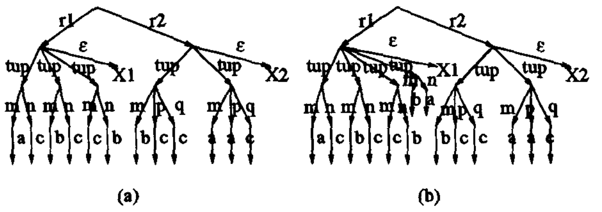


Figure 11: (a) A relational database with markers. (b) Same database after an update.

of [GL95], which we describe next. A **transaction** is a sequence of atomic transactions of the form (1) $DB := DB +\!\!+_{\mathcal{Y}'} \Delta$, for $\mathcal{Y}' \subseteq \mathcal{Y}$, or (2) $DB := DB\, replace_{\mathcal{X}'}\, \Box$. We have defined $+\!\!+$ earlier. For the *replace* operation, assume DB is given in the canonical form (equation 1), and $\mathcal{X}' = \{X_k, X_{k+1}, \ldots, X_m, X_{m+1}, \ldots, X_p\}$ (that is $\mathcal{X} \cap \mathcal{X}' = \{X_k, \ldots, X_m\}$). Then, for $\Box$ a tree with $\mathcal{X}'$ entries, say $\Box = (X_k := \Box_k; X_{k+1} := \Box_{k+1}; \ldots; X_p := \Box_p)$, $DB\; replace_{\mathcal{X}'}\, \Box$ is defined to be:

$$X_1 +\!\!+_{\mathcal{X}'} \mathsf{rec}_{\mathcal{X}'}\ (X_1 := t_1, \ldots, X_{k-1} := t_{k-1}, \\ X_k := \Box_k, \ldots X_p := \Box_p)$$

Note that *replace* is rather unorthodox, in that it is defined on the particular representation of the database, and not on its meaning. More precisely, one can find trees DB and DB' which are bisimilar, hence equal for our purposes, but for which $DB\; replace\, \Box$ is not bisimilar to $DB'\; replace\, \Box$.

Example 5.1 Consider the relational database DB in Figure 6. To model insertions in the relations r_1 and r_2, we redesign DB by introducing two markers X_1 and X_2, as in Figure 11. Then a traditional insert operation $r_1 := r_1 \cup \{(b, a)\}$ is expressed as:

$$DB := DB +\!\!+_{X_1} (X_1 := \{tup \Rightarrow \{m \Rightarrow b, n \Rightarrow a\} \cup X_1\})$$

The result is shown in Figure 11. Note that we explicitly include $\ldots \cup X_1$, in order to allow for future insertions.

6 Main Result

Our key technique for query decomposition and view maintenance is to transform queries into *decomposable* queries.

Definition 6.1 *Let Q be a query in UnQL. We say that Q is* **decomposable** *iff:*

1. *For all t, t', $Q(t +\!\!+ t') = Q(t) +\!\!+ Q(t')$, and*
2. *For all t, t', $Q(t; t') = (Q(t); Q(t'))$.*

t' in item 1 and both t and t' in item 2 are trees with named inputs, hence the definition only makes sense for queries acting on databases with named inputs.

View maintenance for decomposable queries is easy: e.g. after an update $DB' := DB +\!\!+ \Delta$, we have to update the view $V = Q(DB)$ to $V' := V +\!\!+ Q(\Delta)$. Similarly, we show in Section 7 that query decomposition, in the sense of Section 1, is easily done for decomposable queries. For this we need:

Proposition 6.2 *If Q is decomposable then:*

$$Q(\mathsf{rec}\ (X_1 := t_1; \ldots X_k := t_k)) = \\ \mathsf{rec}\ (Q(X_1 := t_1); \ldots Q(X_k := t_k))$$

Proof: We only give an informal argument. Since rec t is the same as the infinite unfolding $t +\!\!+ t +\!\!+ \ldots$, we have $Q(\mathsf{rec}\ t) = Q(t +\!\!+ t +\!\!+ \ldots) = Q(t) +\!\!+ Q(t) +\!\!+ \ldots = \mathsf{rec}\ Q(t)$. Next we apply item 2 of Definition 6.1.

As we shall see, not every UnQL query is decomposable. However, in the theorem below, we prove that every query in the "positive, join-free fragment" of UnQL (definition in Section 8) can be obtained from a decomposable one with minor post-processing: this is the main result of our paper, and we will sketch the proof in Section 9.

Theorem 6.3 *For any query Q in the positive, join-free fragment of UnQL, there exists a (positive, join-free) query Q' and constant tree α such that the following conditions hold:*

1. *$Q = \alpha +\!\!+ Q'$.*
2. *Q' is decomposable.*
3. *The cost of evaluating Q' is no larger that that of a naive evaluation of Q.*

7 Applications

We show here how Theorem 6.3 can be applied to query decomposition and view maintenance.

7.1 Query Decomposition

For illustration we shall assume that the database is stored on two sites s and s': the technique described here generalizes straightforwardly to an arbitrary number of sites. As in Section 4 we assume that DB is given in the canonical form $X_1 +\!\!+ DB_0$ where $DB_0 = \mathsf{rec}_{\mathcal{X}}\ (t; t')$ with t residing on site s and t' on s'. Let Q_0 be the query $Q_0(DB_0) \stackrel{\text{def}}{=} Q(X_1 +\!\!+ DB_0)$[3]. We proceed as follows:

1. Apply Theorem 6.3 to decompose Q_0 into: $Q_0(DB_0) = \alpha +\!\!+ Q'(DB_0)$. Hence we have to compute $Q'(DB_0)$, which is $Q'(\mathsf{rec}_{\mathcal{X}}\ (t; t'))$.
2. By item 2 of Theorem 6.3 and Proposition 6.2, $Q'(\mathsf{rec}_{\mathcal{X}}\ (t; t')) = \mathsf{rec}_{\mathcal{X}'}\ (Q'(t); Q'(t'))$, where $\mathcal{X}'$ is the set of inputs and outputs of $(Q'(t); Q'(t'))$. Hence we compute $Q'(t)$ at site s, and independently $Q'(t')$ at site s'.
3. Send $Q'(t)$ and $Q'(t')$ to the client site: this step is potentially inefficient and will be refined below. The client now holds $(Q'(t); Q'(t'))$.
4. The client computes $\mathsf{rec}_{\mathcal{X}'}\ (Q'(t); Q'(t'))$ (which consists in drawing ε edges from outputs to inputs), thus obtaining $Q'(DB_0)$. Finally it computes $Q_0(DB_0)$, as shown in step 1.

[3]Theorem 6.3 only applies to queries expecting databases with named inputs, hence it does apply to Q_0, but not to Q directly.

Parts of the graphs $Q'(t)$ and $Q'(t')$ sent in step 3 may turn out (during step 4) to be inaccessible from from the root of the final result $Q_0(DB_0)$. To avoid sending these useless fragments in step 3, we refine it as follows:

- Compute the *accessibility graph* G for $(Q'(t); Q'(t'))$: nodes are markers $\mathcal{X}'$, edges are pairs (X, X') s.t. X' is accessible from input $X \in \mathcal{X}'$ in $(Q'(t); Q'(t'))$. The two fragments of G are computed independently on sites s and s', then sent to the client.
- The client computes the transitive closure of G: this gives the accessiblity graph for $Q'(DB_0) = \mathsf{rec}_{\mathcal{X}'}\ (Q'(t); Q'(t'))$. Using this and the tree α, the client computes the set $\mathcal{X}'_0 \subseteq \mathcal{X}'$ of markers acually used in the concatenation $Q_0(DB_0) = \alpha +\!\!+\ Q'(DB_0)$, i.e. $\alpha +\!\!+_{\mathcal{X}'} Q'(DB_0)$ is equal to $\alpha +\!\!+_{\mathcal{X}'_0} Q'(DB_0)$.
- The client sends the set $\mathcal{X}'_0$ to s and s'. Next, step 3 is resumed, but having each server sent only the "accessible" fragment of $Q'(t)$ (and $Q'(t')$ respectively) to the client, i.e. that under the inputs in $\mathcal{X}'_0$.

Finally we argue that the total cost of evaluating our decomposed query is no larger than that of computing $Q(DB)$ on a centralized version of DB, using a naive evaluation method. Indeed Q and Q' are equally expensive to compute. The key observation is that, under a non-optimized evaluation, $Q'(\mathsf{rec}\ (t; t'))$ and $\mathsf{rec}\ (Q(t); Q(t'))$ have the same cost too. An optimized evaluation however might do a better job with $Q'(\mathsf{rec}\ (t; t'))$, because it can avoid processing unaccessible parts.

7.2 View Maintenance

Let a view V be defined by the query Q, $V \overset{\text{def}}{=} Q(DB)$. For simplicity we will assume that DB is a database with named inputs. The view maintenance method consists of the following. Let $DB = X_1 +\!\!+\ DB_0$ be in $\mathit{Tree}_{\mathcal{Y}}$, $Q_0(DB_0) \overset{\text{def}}{=} Q(X_1 +\!\!+\ DB_0)$ be as before, and let $Q_0 = \alpha +\!\!+\ Q'$ be the decomposition of Q_0 given by Theorem 6.3.

1. When the view $V = Q(DB)$ is defined, compute and store $V' = Q'(DB_0)$. In particular $V = \alpha +\!\!+\ V'$, and V' encapsulates the markers $\mathcal{Y}$.
2. Assume that the database is updated as $DB := DB +\!\!+_{\mathcal{Y}'} \Delta$, $\mathcal{Y}' \subseteq \mathcal{Y}$. Send both the set of markers $\mathcal{Y}'$ and the tree Δ from the server to the client[4]. Compute here $Q'(\Delta)$ and update the materialized view to $V' := V' +\!\!+ Q'(\Delta)$. This is correct because Q' is decomposable.
3. Assume now that the database is updated to $DB := DB\ \mathit{replace}_{\mathcal{X}'}\ \Box$. Again $\mathcal{X}'$ and $\Box$ are sent to the client, which updates its view to $V' := V'\ \mathit{replace}\ Q'(\Box)$. Again, this is correct, because Q' is decomposable.

[4]Strictly speaking the set of markers $\mathcal{Y}'$ can be extracted from Δ: it is exactly the set of inputs of Δ.

Note that the new view can be computed only from the old one and from the increment Δ or $\Box$: it does not depend directly on DB. This is possible because Q is join-free.

8 The calculus: UnCAL$_0$

[BDHS96a] introduces the language UnQL as a query language for browsing unstructured data, or data whose structure is only partially known. It is shown that UnQL is equivalent to a calculus, called UnCAL, much in the same way in which a certain fragment of SQL is equivalent to the relational algebra. Here we briefly review UnCAL. This is a rather dry formalism, and we refer the reader to [BDHS96a] for a presentation of the friendlier, and equivalent, UnQL language.

Figure 12 lists the constructs of a monotone fragment of UnCAL, which we call UnCAL$_0$. UnCAL$_0$ is obtained by restricting UnCAL in two ways: (1) the emptiness test *isEmpty?*(t) is removed, and (2) the *gext* construct is replaced with the more restrictive *vext* one. One can show that all queries expressed in UnCAL$_0$ are "monotone", in a way which can be made formal [Bun95].

We briefly describe the constructs in Figure 12, where Q, Q_1, Q_2 are queries, a, a' are label variables or constants, p is a user-defined predicate. DB stands for the input database, $\{\}$ returns the empty set (one node, no edges), $\{a \Rightarrow Q\}$ returns a singleton tree (here a is a label constant or a label variable), while $Q_1 \cup Q_2$ returns the union of two trees. We can represent the latter conveniently using ε edges, as in Figure 12. $Q_1 +\!\!+_{\mathcal{A}} Q_2$ concatenates Q_1 with Q_2 (Section 3).

The most complex construct is $\mathit{vext}_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$. Here Q_1 is an expression denoting a tree with inputs $\mathcal{A}$ and outputs $\mathcal{A}$, which may have x as a free label variable. Then $\mathit{vext}_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$ returns a tree obtained from Q_2 by replacing every edge labeled a with the tree $Q_1[a/x]$. Figure 12 tries to illustrate this action for the case when $\mathcal{A} = \{A\}$ has a single marker. We will illustrate *vext* below.

Next, $A := Q$ renames the unique, anonymous input of Q to A: i.e. when $Q \in \mathit{Tree}_{\mathcal{X}}$, then $A := Q$ is in $\mathit{Tree}^{\{A\}}_{\mathcal{X}}$. In the juxtaposition $(Q_1; Q_2)$, the inputs of Q_1 and Q_2 have to be disjoint: the result will be the union of the two graphs (which is *not* $Q_1 \cup Q_2$). That is, when $Q_1 \in \mathit{Tree}^{\mathcal{X}_1}_{\mathcal{Y}}, Q_2 \in \mathit{Tree}^{\mathcal{X}_2}_{\mathcal{Y}}$ and $\mathcal{X}_1 \cap \mathcal{X}_2 = \emptyset$, then $(Q_1; Q_2)$ is in $\mathit{Tree}^{\mathcal{X}_1 \cup \mathcal{X}_2}_{\mathcal{Y}}$. We abbreviate $(Q_1; (Q_2; (\ldots; Q_n \ldots)))$ with $(Q_1; \ldots; Q_k)$. $()$ is the null tree (no nodes, no edges). We have $(Q_1; Q_2) = (Q_2; Q_1)$ and $(Q; ()) = Q$. A stands for a marker.

A *query* is an expression with no free variables, except for DB, which is the query input. A query may use markers of its own, in constructs like $+\!\!+_{\mathcal{A}}, \mathit{vext}_{\mathcal{A}}, A := Q, A$, which we denote with $A, A', A_1, A_2, \ldots$, and which are distinct from the markers $X, X_1, \ldots, Y, Y_1, \ldots$ in the database DB. This

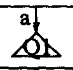
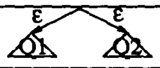
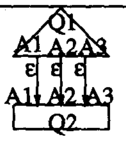
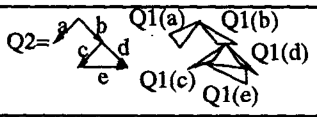
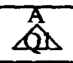
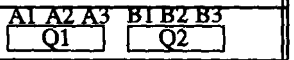
Query Q	Description	Meaning
DB	input database	
$\{\}$	the empty set	
$\{a \Rightarrow Q_1\}$	singleton with edge label a	
$Q_1 \cup Q_2$	union	
$Q_1 +\!\!+_{\mathcal{A}} Q_2$	concatenation	
$vext_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$	expansion	
if C then Q_1 else Q_2	conditional	C a condition of the form $a = a'$ or $p(a)$
$A := Q_1$	single input named A	
$(Q_1; Q_2)$	juxtaposition	
$()$	the null tree	
A	marker A	

Figure 12: The UnCAL$_0$ calculus.

is much like in the case of oid's: $X, X_1, \ldots$ are like external oid's, and the query Q cannot access them. We allow one exception to this rule: a query on a database with named inputs may use the unique marker X_1 denoting the "root" input to DB: thus, the query $Q(DB) \stackrel{\text{def}}{=} X_1 +\!\!+_{X_1} DB$ is legal, but $X_2 +\!\!+_{X_2} DB$ or $DB +\!\!+_{Y_1} (Y_1 := \{a\})$ are not.

The main construct of the language is *vext*. First we will explain how $vext_{\mathcal{A}}(\lambda x.Q)(Q')$ interacts with the input and output markers of Q'. Recall that Q must have $\mathcal{A} = \{A_1, \ldots, A_p\}$ as inputs and outputs, and assume that Q' has inputs $\mathcal{X} = \{X_1, \ldots, X_m\}$ and outputs $\mathcal{Y} = \{Y_1, \ldots, Y_n\}$. By definition *vext* returns a tree with $p \cdot m$ inputs and $p \cdot n$ outputs. To name them, we denote with $U \cdot V$ a new, distinct marker, for any two markers U, V. We require this operation to be injective, i.e. $U \cdot V = U' \cdot V' \Longrightarrow U = U'$ and $V = V'$. Then, by definition, the inputs of $vext_{\mathcal{A}}(\lambda x.Q)(Q')$ are $\mathcal{A} \cdot \mathcal{X} \stackrel{\text{def}}{=} \{A_i \cdot X_j \mid A_i \in \mathcal{A}, X_j \in \mathcal{X}\}$, and the outputs are $\mathcal{A} \cdot \mathcal{Y}$. By convention, when Q' has a single anonymous input, then $vext_{\mathcal{A}}(\lambda x.Q)(Q')$ has inputs $\mathcal{A}$. We illustrate next *vext* with a detailed example, and refer to [BDHS96a] for a general definition of *vext*.

Example 8.1 Recall the query in Example 1.1 which returns all papers in the Computer Science Department. We consider here a variant, which stops following a path once it finds an *"Papers"* edge before a *"CS-Department"*: intuitively this is an optimization, if one assumes that no department edge ever occurs after a *"Papers"* edge. Then Q can be expressed as follows: $A_1 +\!\!+_{\mathcal{A}} (vext_{\mathcal{A}}(\lambda x.Q_1(x))(DB))$, where $\mathcal{A} = \{A_1, A_2, A_3\}$ and $Q_1(x)$ is[5]:

if $x =$ *"CS-Department"* then
 $(A_1 := A_2; A_2 := A_2; A_3 := \{x \Rightarrow A_3\})$
else if $x =$ *"Papers"* then
 $(A_1 := \{\}; A_2 := \{x \Rightarrow A_3\}; A_3 := \{x \Rightarrow A_3\})$
else $(A_1 := A_1; A_2 := A_2; A_3 := \{x \Rightarrow A_3\})$

$Q_1(x)$ is best visualized graphically, as in Figure 13 (a). Figure 13 (b) contains an example of a database with three outputs $\mathcal{X} = \{X, Y, Z\}$. Then $vext_{\mathcal{A}}(\lambda x.Q_1)(DB)$ returns a database with 9 outputs $\mathcal{A} \cdot \mathcal{X}$, which is represented in Figure 13 (c). After eliminating ϵ edges and the unaccessible part of the output graph we obtain the database in Figure 13 (d). Note that only the outputs $A_3 \cdot Y$ and $A_2 \cdot Z$ remain in the result. The intuition is that the markers X, Y, Z are "copied" in the output, and "tagged" with the region where they were found, as described in Section 2: that is $A_3 \cdot Y$ means that Y has been found in region 3, while $A_2 \cdot Z$ means that Z was found in region 2. Note that marker X disappears: no matter how we extend DB at marker X (e.g. an update, or a link to another site), $Q(DB)$ will not be affected.

Q is not decomposable in the sense of Definition 6.1. However the sub-query $Q'(t) = vext_{\mathcal{A}}(\lambda x.Q_1)(t)$ *is* decomposable: when $t \in Tree^{\mathcal{X}}_{\mathcal{Y}}, t' \in Tree^{\mathcal{Y}}_{\mathcal{Z}}$ then $Q'(t) \in Tree^{\mathcal{A}\cdot\mathcal{X}}_{\mathcal{A}\cdot\mathcal{Y}}$, $Q'(t') \in Tree^{\mathcal{A}\cdot\mathcal{Y}}_{\mathcal{A}\cdot\mathcal{Z}}$, and $Q'(t +\!\!+_{\mathcal{Y}} t') =$

[5]The original query in Example 1.1 would have had $A_1 := A_1$ instead of $A_1 := \{\}$ in the case $x =$ *"Papers"*.

$Q'(t) +\!\!+_{A \cdot y} Q'(t')$. Moreover, Q is easily recovered from Q' as: $Q = (A_1 +\!\!+_{A_1} Q')$. Theorem 6.3 generalizes this observation.

The **query evaluation** method for UnCAL consists in manipulating labeled graphs, as suggested in Figure 12, and occasionally "cleaning up" the graph, by reducing it under bisimilarity (this also eliminates the unaccessible parts). For the main construct, *vext*, there are two evaluation strategies, which we illustrate on $vext(\lambda x.Q)(DB)$. (1) Start from the root in DB and traverse it recursively, using Q as an automata. For the example in Figure 13 (b), this will produce directly the graph in (d), without ever touching the unaccessible parts. (2) Process each edge of DB independently, by replacing it with Q, a process illustrated in Figure 13 (c): here a clean-up phase is required to get to (d). Method (1) can be more efficient, and we called the "optimized" evaluation; it needs special care to deal with cycles; method (2) makes more sense on databases with more than one input, since there is no root to start from, but is more "naive" because it may end up constructing unnecessary fragments of a graph, which are later eliminated. The statement in item 3 of Theorem 6.3 holds for both methods. However in the query decomposition method recall that we also need to replace $Q(\mathsf{rec}\ (t; t'))$ with $\mathsf{rec}\ (Q(t); Q(t'))$ for some decomposable Q: the cost of evaluating both of them is the same under method (2), but method (1) could be more efficient on $Q(\mathsf{rec}\ (t; t'))$, because it may avoid accessing parts of the database which are not needed. Therefore our query decomposition method has the same cost as computing $Q(DB)$ on a centralized version of DB using method (2), but may be more expensive than $Q(DB)$, when method (1) is used.

We call a UnCAL$_0$ expression Q *constant* if it doesn't mention the input database DB. Note that in the example above, the sub-query Q_1 in $vext_{\mathcal{A}}(\lambda x.Q_1)(DB)$ is a constant expression. We call a query Q *join-free* if in every sub-query of the form $vext_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$, (1) Q_1 is constant and (2) Q_1 does not use any additional output markers besides $\mathcal{A}$. It is easy to extend the results in [BDHS96a] to show that the relational algebra operations union, selection and projection over relational databases expressed as trees, can be expressed as join-free UnCAL$_0$ queries, and that intersection, cartesian product, and eqjoin can be expressed in UnCAL$_0$ (but not as join-free queries). Condition (2) is imposed because we want to disallow self concatenation, like in $DB +\!\!+_B DB$, in order for Theorem 6.3 to hold. Direct self concatenation is not possible in UnCAL$_0$, because we cannot use DB's output markers in a query. But without rule (2) we can still do a self-concatenation in the form $vext_{\mathcal{A}}(\lambda x.Q_1)(DB) +\!\!+_B DB$ in which Q_1 uses the marker $B \notin \mathcal{A}$ to place it deep into DB and leave it there.

Finally, for the purpose of the results in Section 6, we call an UnQL query Q ***positive*** iff its translation into UnCAL given in [BDHS96a] is in UnCAL$_0$ (i.e. doesn't use ***isEmpty?***). Similarly, we call Q ***join-free*** iff its translation is join-free. It is not difficult to design syntactic conditions which are sufficient for UnQL queries to be positive and/or join-free.

9 Decomposition Rules

We sketch here the proof of Theorem 6.3, by describing the decomposition rules which, when applied to some query Q, transform it into $\alpha +\!\!+ Q'$. The rules are given in Figure 14, and apply inductively on the sub-queries of Q. Note that in $vext_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$, Q_1 is a constant expression and hence does not need to be transformed. As a consequence the inductive transformation rules never reach an $\mathsf{if} - \mathsf{then} - \mathsf{else}$ construct. Before applying the rules, we transform the query Q ensuring that distinct $vext_{\mathcal{A}}$ subexpressions in Q use disjoint sets of markers $\mathcal{A}$: if not, then we simply rename the markers in some of the $vext_{\mathcal{A}}$ constructs.

For the base case, when Q is $DB \in Tree^{\mathcal{X}}_{\mathcal{Y}}$, α has to be taken as the "identity on the inputs $\mathcal{X}$", i.e. $(X_1 := X_1; \ldots; X_m := X_m)$. This is not expressible in UnCAL$_0$, because we do not have access to the sets of markers $\mathcal{X}$, so we add a new construct, $id_{\mathcal{X}}$, denoting the identity on the inputs of DB. Hence, strictly speaking α will be an expression in UnCAL$_0$ + $id_{\mathcal{X}}$.

First we explain the juxtapositions in Figure 14. Consider any two sub-queries Q_1 and Q_2 of some larger query Q, and let Q'_1 and Q'_2 be their transformations. Q'_1 and Q'_2 may have common input markers, hence the $(Q'_1; Q'_2)$ construct is not quite correct. However we observe that the only possible common inputs are those in DB. More precisely, for Q'_1 (and similarly Q'_2) one of two cases arise: (1) every input in Q'_1 is of the form $A \cdot (\ldots)$, i.e. tagged with some marker A used in a $gext_{\mathcal{A}}$ subexpression of Q_1, hence private to Q_1; in this case all inputs in Q'_1 are disjoint from those in Q'_2 and the juxtaposition $(Q'_1; Q'_2)$ is correct. (2) $Q'_1 = (DB; Q''_1)$, with all inputs of Q''_1 tagged with private marker of Q_1. In this case we inspect Q'_2: if it also has the form $(DB; Q''_2)$, then we adopt for $(Q'_1; Q'_2)$ the meaning $(DB; Q''_1; Q''_2)$, i.e. copy only once the common part which is DB.

We illustrate the correctness of the rules in Figure 14 for *vext*, $\cup$ and $+\!\!+_{\mathcal{A}}$. Consider the query $Q = vext_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$. First apply induction hypothesis to get $Q_2 = \alpha_2 +\!\!+ Q'_2$ (recall that Q_1 is constant). Then Figure 14 defines $Q' \overset{\text{def}}{=} vext_{\mathcal{A}}(\lambda x.Q_1)(Q'_2)$ and $\alpha \overset{\text{def}}{=} vext_{\mathcal{A}}(\lambda x.Q_1)(\alpha_2)$. We will use the following two equations:

$$vext_{\mathcal{A}}(\lambda x.Q_1)(t_1 +\!\!+ t_2) = vext_{\mathcal{A}}(\lambda x.Q_1)(t_1) +\!\!+ vext_{\mathcal{A}}(\lambda x.Q_1)(t_2) \quad (2)$$

$$vext_{\mathcal{A}}(\lambda x.Q_1)(t_1; t_2) = (vext_{\mathcal{A}}(\lambda x.Q_1)(t_1); vext_{\mathcal{A}}(\lambda x.Q_1)(t_2)) \quad (3)$$

Equation 2 is from [BDHS96a], while 3 is a direct consequence of the definition of *vext*. First we check that $Q = \alpha +\!\!+ Q'$, which follows from equation 2. Next, using both equations 2 and 3, and the induction hypothesis that Q'_2 is decomposable, one can easily check that Q' is decomposable too.

Consider now the case $Q = Q_1 \cup Q_2$. First apply induction hypothesis to get $Q_i = \alpha_i +\!\!+ Q'_i$, $i = 1, 2$.

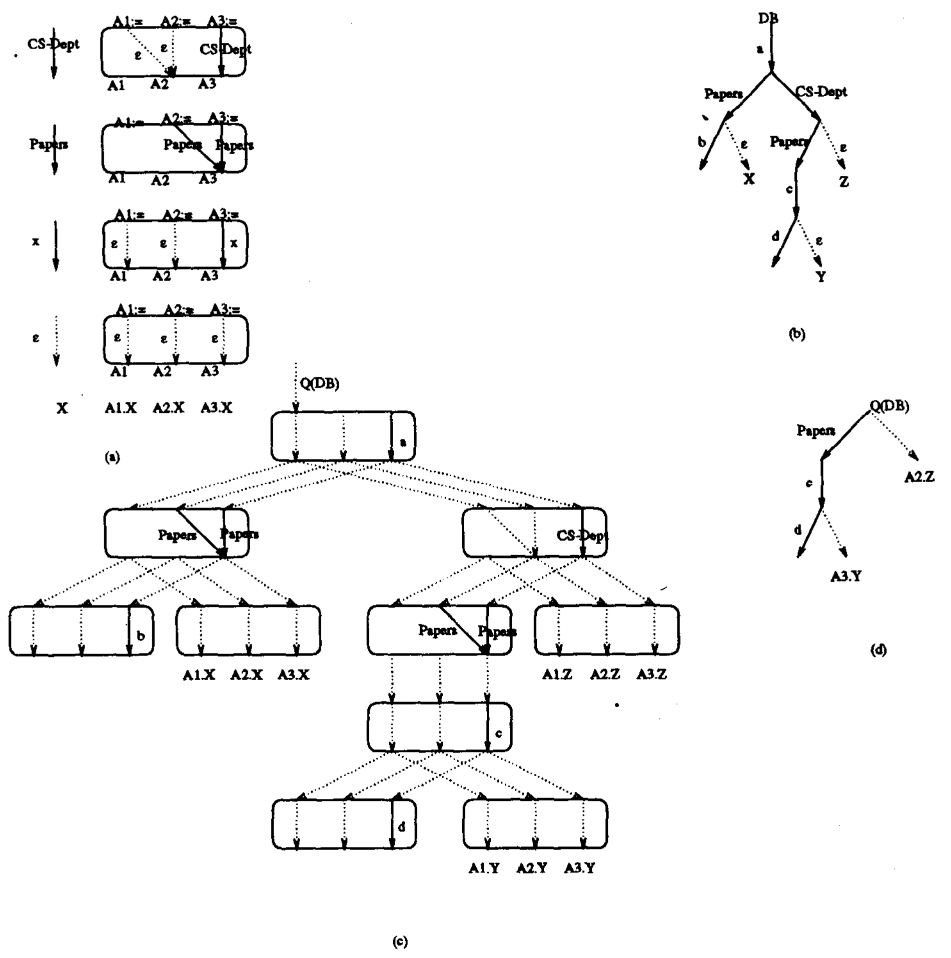


Figure 13: Illustration of $vext(\lambda x.Q)(DB)$.

Then Figure 14 defines $Q' \overset{\text{def}}{=} (Q'_1; Q'_2)$, and $\alpha \overset{\text{def}}{=} \alpha_1 \cup \alpha_2$. First check item 1 of Theorem 6.3:

$$\begin{aligned} & \alpha ++ Q' = \\ = \;& (\alpha_1 \cup \alpha_2) ++ (Q'_1; Q'_2) \\ = \;& (\alpha_1 ++ Q'_1) \cup \\ & (\alpha_2 ++ Q'_2) = \\ = \;& Q_1 \cup Q_2 \end{aligned}$$

The first equality is true because of the separation of the inputs markers for Q'_1 and Q'_2, as explained above. They may share only input markers from the inputs of DB, and in this case their common part is exactly DB: hence $(\alpha_1 \cup \alpha_2) ++ (Q'_1; Q'_2) = (\alpha_1 ++ Q'_1) \cup (\alpha_2 ++ Q'_2)$.

Next we check item 2:

$$\begin{aligned} & Q'(t_1 ++ t_2) = \\ = \;& (Q'_1(t_1 ++ t_2); Q'_2(t_1 ++ t_2)) = \\ = \;& (Q'_1(t_1) ++ Q'_1(t_2); \\ & Q'_2(t_1) ++ Q'_2(t_2)) = \\ = \;& (Q'_1(t_1); Q'_2(t_1)) ++ \\ & (Q'_1(t_2); Q'_2(t_2)) = \\ = \;& Q'(t_1) ++ Q'(t_2) \\ & Q'(t_1; t_2) = \\ = \;& (Q'_1(t_1; t_2); Q'_2(t_1; t_2)) = \\ = \;& (Q'_1(t_1); Q'_2(t_1); Q'_1(t_2); Q'_2(t_2)) = \\ = \;& (Q'_1(t_1); Q'_1(t_2); Q'_2(t_1); Q'_2(t_2)) = \\ = \;& (Q'(t_1); Q'(t_2)) \end{aligned}$$

For the $Q = Q_1 ++_{\mathcal{A}} Q_2$ construct, we again apply induction first, $Q_i = \alpha_i ++ Q'_i$, $i = 1, 2$. Obviously $(Q'_1; Q'_2)$ is decomposable (as above), so it remains to show that $Q_1 ++_{\mathcal{A}} Q_2 = (\alpha_1 ++_{\mathcal{A}} \alpha_2) ++ (Q'_1; Q'_2)$. For this, we first show by induction that every output marker of Q'_1 is "tagged" with some output marker in $\mathcal{Y}$. Hence none of markers in $\mathcal{A}$ may appear in Q'_1 (this is a consequence of condition (2) of the join-freeness definition of Section 8). Hence we have:

$$\begin{aligned} & Q_1 ++_{\mathcal{A}} Q_2 = \\ = \;& (\alpha_1 ++ Q'_1) ++_{\mathcal{A}} (\alpha_2 ++ Q'_2) \\ = \;& \alpha_1 ++_{\mathcal{A}} (Q'_1; \alpha_2) ++ Q'_2 \\ = \;& (\alpha_1 ++_{\mathcal{A}} \alpha_2) ++ Q'_1 ++ Q'_2 \\ = \;& (\alpha_1 ++_{\mathcal{A}} \alpha_2) ++ (Q'_1; Q'_2) \end{aligned}$$

10 Conclusions

We have described a query decomposition method which can be used to compute efficiently queries

Query Q	Q'	α
DB	DB	$id_{\mathcal{X}}$
$\{\}$	$()$	$\{\}$
$\{a \Rightarrow Q_1\}$	Q'_1	$\{a \Rightarrow \alpha_1\}$
$Q_1 \cup Q_2$	$(Q'_1; Q'_2)$	$\alpha_1 \cup \alpha_2$
$Q_1 +\!\!+_{\mathcal{A}} Q_2$	$(Q'_1; Q'_2)$	$\alpha_1 +\!\!+_{\mathcal{A}} \alpha_2$
$vext_{\mathcal{A}}(\lambda x.Q_1)(Q_2)$	$vext_{\mathcal{A}}(\lambda x.Q_1)(Q'_2)$	$vext_{\mathcal{A}}(\lambda x.Q_1)(\alpha_2)$
if C then Q_1 else Q_2	N/A	N/A
$A := Q_1$	Q'_1	$A := \alpha_1$
$(Q_1; Q_2)$	$(Q'_1; Q'_2)$	$(\alpha_1; \alpha_2)$
$()$	$()$	$()$
A	$()$	A

Figure 14: Rules for transforming any UnCAL$_0$ join-free query Q into $\alpha +\!\!+ Q'$, with Q' decomposable.

in unstructured query languages on distributed data sources. The decomposition is "efficient" in the sense that, under a certain naive evaluation strategy, the decomposed query is no more expensive to compute on the distributed database, than were the original if applied to a centralized database. Also, we have proposed two update operations for the underlying data model, and shown that the same query decomposition method can be used to derive a simple view maintenance method.

Our methods work without any knowledge about the structure of the database. While sometimes this can be useful, in general it is more a limitation than a virtue: some simple knowledge about the database could improve the query decomposition. In future work we plan to develop techniques to incorporate such information in the query decomposition.

Acknowledgments I would like to thank Peter Buneman and Gerd Hillebrand for their comments.

References

[BD96] Thomas Ball and Fred Douglis. An internet difference engine and its applications. In *Procs. of COMPCON*, February 1996.

[BDHS96a] Peter Buneman, Susan Davidson, Gerd Hillebrand, and Dan Suciu. A query language and optimization techniques for unstructured data. In *SIGMOD*, 1996.

[BDHS96b] Peter Buneman, Susan Davidson, Gerd Hillebrand, and Dan Suciu. A query language and optimization techniques for unstructured data. Technical Report 96-09, University of Pennsylvania, Computer and Information Science Department, February 1996.

[BDS95] Peter Buneman, Susan Davidson, and Dan Suciu. Programming constructs for unstructured data. In *Proceedings of DBPL'95*, Gubbio, Italy, September 1995.

[BK90] C. Beeri and Y. Kornatzky. A logical query language for hypoertext systems. In *Hypertext: Concepts, Systems, and Applications. Procs. of the European Conf. on Hypertext*, pages 67–80, 1990.

[Bun95] Peter Buneman, 1995. Private communication.

[Cou83] B. Courcelle. Fundamental properties of infinite trees. *Theoretical Computer Science*, 25:95–169, 1983.

[DB96] Fred Douglis and Thomas Ball. Tracking and viewing changes on the web. In *Procs of USENIX Technical Conference*, January 1996.

[GL95] Timothy Griffin and Leonid Libkin. Incremental mainenance of views with duplicates. In *International Conference on Management of Data*, pages 328–339, San Jose, California, June 1995.

[KS95] David Konopnicki and Oded Shmueli. Draft of W3QS: a query system for the World-Wide Web. In *Proc. of VLDB*, 1995.

[LMSS96] Alon Levy, Alberto Mendelzon, Yehoshua Sagiv, and Divesh Srivastava. Answering queries using views. In *Proceedings of the 14th Symposium on Principles of Database Systems*, San Jose, CA, June 1996.

[LRU96] Alon Levy, Anand Rajaraman, and Jeffrey Ullman. Answering queries using limited external processors. In *PODS*, 1996. To appear.

[PGMW95] Y. Papakonstantinou, H. Garcia-Molina, and J. Widom. Object exchange across heterogeneous information sources. In *IEEE International Conference on Data Engineering*, March 1995.

[QRS$^+$95] D. Quass, A. Rajaraman, Y. Sagiv, J. Ullman, and J. Widom. Querying semistructure heterogeneous information. In *International Conference on Deductive and Object Oriented Databases*, 1995.