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Efficient Query Processing for Dynamically Changing Datasets

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ABSTRACT
The ability to efficiently analyze changing data is a key requirement of many real-time analytics applications. Traditional approaches to this problem were developed around the notion of Incremental View Maintenance (IVM), and are based either on the materialization of subresults (to avoid their recomputation) or on the recomputation of subresults (to avoid the space overhead of materialization). Both techniques are suboptimal: instead of materializing results and subresults, one may also maintain a data structure that supports efficient maintenance under updates and from which the full query result can quickly be enumerated. In two previous articles, we have presented algorithms for dynamically evaluating queries that are easy to implement, efficient, and can be naturally extended to evaluate queries from a wide range of application domains. In this paper, we discuss our algorithm and its complexity, explaining the main components behind its efficiency. Finally, we show experiments that compare our algorithm to a state-of-the-art (Higher-order) IVM engine, as well as to a prominent complex event recognition engine. Our approach outperforms the competitor systems by up to two orders of magnitude in processing time, and one order in memory consumption.

1 Introduction
The ability to efficiently analyze changing data is a key requirement of many real-time analytics applications like Stream Processing [20], Complex Event Recognition [9], Business Intelligence [17], and Machine Learning [22].

In this context, we tackle the problem of dynamic query evaluation, where a given query \(Q\) has to be evaluated against a database that is constantly changing. Concretely, when database \(db\) is updated to database \(db + u\) under update \(u\), the objective is to efficiently compute \(Q(db + u)\), taking into consideration that \(Q(db)\) was already evaluated and re-computations could be avoided.

Dynamic query evaluation is of utmost importance if response time requirements for queries under concurrent data updates have to be met or if data volumes are so large that full re-evaluation of queries based on raw data is prohibitive.

The following example illustrates our setting. Assume that we wish to detect potential credit card fraud. Credit card transactions specify their timestamp \((ts)\), account number \((acc)\), and amount \((amnt)\). A typical fraud pattern is that, in a short period of time, a criminal tests a stolen credit card with a few small purchases to then make larger purchases (cf. [18]). Assuming that the short period of time is 1 hour, this pattern could be detected by dynamically evaluating the query in Figure 1. Queries like this may exhibit arbitrary local predicates and multi-way joins with equality as well as inequality predicates. Traditional techniques to process such queries dynamically can be categorized in two approaches: relational and automaton-based. We outline the core principles of relational approaches in the following and refer to [13] for an in-depth discussion of the drawbacks of automaton-based approaches.

Relational approaches such as [2, 10, 16] build upon the technique of Incremental View Maintenance (IVM) [7]. To process a query \(Q\) over a database \(db\), IVM techniques materialize the output \(Q(db)\) and evaluate \(delta\) queries [10]. Upon update \(u\), delta queries use \(db, u\), and the materialized \(Q(db)\) to compute the set of tuples to add/delete from \(Q(db)\) in order to obtain \(Q(db + u)\). If \(u\) is small with respect to the database \(db\), this is expected to be faster than recomputing \(Q(db + u)\) from scratch. To further speed up dynamic query processing, also the result of some subqueries of \(Q\) may be redundantly materialized. This approach is known as Higher-Order IVM (HIVM) [15, 16].

Unfortunately, (H)IVM shows a serious drawback in terms of additional memory overhead, which quickly becomes prohibitive for interactive data analytics scenarios: materialization of \(Q(db)\) requires \(\Omega(|Q(db)|)\) space, where \(|db|\) denotes the size of \(db\). Therefore, when \(Q(db)\) is large, which is often the case in data preparation scenarios for training statistical models, materializing \(Q(db)\) quickly becomes impractical, especially for main-memory based systems. HIVM is even more affected by this problem than IVM since it not only

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*This paper surveys the results of two papers: “The Dynamic Yan- nakakis Algorithm: Compact and Efficient Query Processing Under Updates” [12], which was published in SIGMOD Conference 2017 and “Conjunctive Queries with Inequalities Under Updates” [13], published in PVLDB 11(7), 2018. During the work described, Martin Ugarte was affiliated to ULB and Hannes Voight to TU Dresden.

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materializes the result of \( Q \) but also the results of some subqueries. In fact some of these subresults can be partial join results, which can be larger than both \( db \) and \( Q(db) \). For example, in our fraud query, H)IVM would materialize the results of the join in the shaded area in Figure 1. Intuitively, this join builds the table of all pairs of small transactions that could be part of a credit card fraud if a third relevant transaction occurs. Therefore, if we assume that there are \( N \) small transactions in the time window, all of the same account, this materialization will take \( \Theta(N^2) \) space. This naturally becomes impractical when \( N \) grows.

While these problems are inherent to H)IVM methods, they can be avoided by taking a different approach to dynamic query evaluation: instead of materializing \( Q(db) \) we can build a succinct data structure that (1) supports efficient maintenance under updates and (2) represents \( Q(db) \) in the sense that from it we can generate \( Q(db) \) as efficiently as if it were materialized. In particular, the representation is equipped with index structures so that we can enumerate \( Q(db) \) with constant delay [19]: one tuple at a time, while spending only a constant amount of work to produce each new tuple. This makes the enumeration competitive with enumeration from materialized query results.

In essence, we hence separate dynamic query processing into two stages: (1) an update stage where we only maintain the (small) information that is necessary for result enumeration and (2) an enumeration stage where the query result is efficiently enumerated.

In our work, which is documented in detail in [12] and [13], we are concerned with designing a practical family of algorithms for dynamic query evaluation based on this idea for queries featuring both equi-joins and inequality joins, as well as certain forms of aggregation. Our main insight is that, for acyclic conjunctive queries, such algorithms can naturally be obtained by modifying Yannakakis' seminal algorithm for processing acyclic joins in the static setting [25].

In a first step, we address the problem of efficiently evaluating acyclic aggregate-join queries by providing the Dynamic Yannakakis Algorithm (Dyn) [12]. The representation of query results that underlies this algorithm has several desirable properties:

- \((P_1)\) It allows to enumerate \( Q(db) \) with constant delay.
- \((P_2)\) For any tuple \( \tilde{t} \), it can be used to check whether \( \tilde{t} \in Q(db) \) in constant time.
- \((P_3)\) It requires only \( \mathcal{O}(|db|) \) space and is hence independent of the size of \( Q(db) \).
- \((P_4)\) It features efficient maintenance under updates: given update \( u \) to \( db \), we can update the representation of \( Q(db) \) to a representation of \( Q(db + u) \) in time \( \mathcal{O}(|db| + |u|) \). In contrast, (H)IVM may require \( \Omega(|u| + |Q(db + u)|) \) time in the worst case. For the subclass of \( q \)-hierarchical (queries [4], our update time is \( \mathcal{O}(|u|) \).

Based on this technique to dynamically process queries with equi-joins, we provide the core intuition of a generalization of the Dynamic Yannakakis Algorithm to conjunctive queries with arbitrary \( \theta \)-joins. We show that, in the specific case of inequality joins, this generalization improves the state of the art for dynamically processing inequality joins by performing consistently better, with up to two orders of magnitude improvements in processing time and one order in memory consumption.

It is important to note that we consider query evaluation in main memory and measure time and space under data complexity [21]. That is, the query is considered to be fixed and not part of the input. This makes sense under dynamic query evaluation, where the query is known in advance and the data is constantly changing.

## 2 Preliminaries

### Query Language

Throughout the paper, let \( x, y, z, \ldots \) denote variables (also commonly called column names or attributes). A hyperedge is a finite set of variables. We use \( \pi, \ldots \) to denote hyperedges. A Generalized Conjunctive Query (GCQ) is an expression of the form

\[
Q = \pi \left( r_1(\sigma_{r_1}) \land \cdots \land r_n(\sigma_{r_n}) \right) \land \bigwedge_{i=1}^m \theta_i(\sigma_{r_i}).
\]

Here \( r_1, \ldots, r_n \) are relation symbols; \( \sigma_{r_1}, \ldots, \sigma_{r_n} \) are hyperedges (of the same arity as \( r_1, \ldots, r_n \)); \( \theta_1, \ldots, \theta_m \) are predicates over \( \sigma_{r_1}, \ldots, \sigma_{r_n} \), respectively; and both \( \pi \) and \( \bigcup_{i=1}^m \sigma_{r_i} \) are subsets of \( \bigcup_{i=1}^m \sigma_{r_i} \). We treat predicates abstractly: for our purpose, a predicate over \( \pi \) is a (not necessarily finite) decidable set \( \theta \) of tuples over \( \pi \). For example, \( \theta(x, y) = x < y \) is the set of all tuples \((a, b)\) satisfying \( a < b \). We indicate that \( \theta \) is a predicate over \( \pi \) by writing \( \theta(\pi) \). Throughout the paper, we consider only non-nullary predicates with \( \pi \neq \emptyset \).

### Example 2.1.

The following query is a GCQ.

\[
\pi_{y,z,w,u}(r(x, y) \land s(y, z, w) \land t(u, v)) \mid x < z \land w < u
\]

Intuitively, the query asks to take the natural join of \( r(x, y) \) and \( s(y, z, w) \), form the cartesian product with \( t(u, v) \), and subsequently select those tuples that satisfy \( x < z \) and \( w < u \).

We call the output variables of \( Q \) and denote them \( out(Q) \). If \( \pi = \pi_1 \cup \cdots \cup \pi_m \) then \( Q \) is called full and we may omit the symbol \( \pi \) for brevity. We denote by \( full(Q) \) the full GCQ obtained from \( Q \) by setting \( out(Q) \) to \( \pi_1 \cup \cdots \cup \pi_m \). The elements \( r_i(\sigma_{r_i}) \) are called atoms. \( at(Q) \) denotes the set of all atoms in \( Q \), and \( pred(Q) \) the set of all predicates in \( Q \). A conjunctive query (or CQ) is a GCQ where \( pred(Q) = \emptyset \).

### Semantics

We evaluate GCQs over Generalized Multiset Relations (GMRs for short) [12, 15, 16]. Let \( dom(x) \) denote the domain of variable \( x \). As usual, a tuple over \( \pi \) is a mapping \( \tilde{t} \) that assigns a value from \( dom(x) \) to every \( x \in \pi \). A GMR \( R \) over \( \pi \) is a function \( R : [\pi] \rightarrow \mathbb{Z} \) mapping tuples over \( \pi \) to integers such that \( R(\tilde{t}) \neq 0 \) for finitely many tuples \( \tilde{t} \). Here, \( [\pi] \) denotes the set of all tuples over \( x \). In contrast to classical multisets, the multiplicity of a tuple in a GMR can
R is a GMR over \( \bar{\mathcal{R}} \) and \( I \) is an index of \( R \) on \( \bar{\mathcal{R}} \subseteq \bar{\mathcal{R}} \) then we assume that for every \( \bar{\mathcal{R}} \)-tuple \( \bar{s} \) we can retrieve in \( O(1) \) time a pointer to the GMR \( R \Join \bar{s} \), which is the GMR over \( \bar{\mathcal{R}} \) consisting of all tuples that project to \( \bar{s} \):

\[
R \Join \bar{s} \in \text{GMR}[\bar{\mathcal{R}}] : \bar{t} \mapsto \begin{cases} R(\bar{t}) & \text{if } \bar{t}[\bar{y}] = \bar{s} \\ 0 & \text{otherwise} \end{cases}
\]

Moreover, we assume that single-tuple insertions and deletions to \( R \) can be reflected in the index in \( O(1) \) time and that an index takes space linear in |\( R \)|. Essentially, our assumptions amount to perfect hashing of linear size [8]. Although this does not directly match a realistic setting, it is well known that complexity results for this model can be translated, through amortized analysis, to average complexity in real-life implementations [8].

### 3 Dynamic Yannakakis

In this section we formulate \( \text{DY} \), a dynamic version of the Yannakakis algorithm [23], that focuses on the evaluation of CQs. How to deal with \( \theta \)-joins is discussed in Section 4.

#### 3.1 Intuition

A data structure \( D \) supports enumeration of a set \( E \) if there is a routine \( \text{enum}(D) \) outputs each element of \( E \) exactly once. Such enumeration occurs with delay \( d \) if the time until the first output; the time between any two consecutive outputs; and the time between the last output and the termination of \( \text{enum}(D) \), are all bounded by \( d \). \( D \) supports enumeration of a GMR \( R \) if it supports enumeration of the set \( \text{En} \{ (\bar{t}, R(\bar{t})) \mid \bar{t} \in \text{ supp}(R) \} \). When evaluating a GCQ \( Q \) over a database \( db \), we will be interested in representing the elements of \( Q(db) \) by means of a data structure \( D_{db} \), such that we can enumerate \( Q(db) \) from \( D_{db} \). If, for every \( \bar{d} \), the delay to enumerate \( Q(db) \) from \( D_{db} \) is independent of \( \bar{d} \) then we say that the enumeration occurs with constant delay [19].

As a trivial example of constant delay enumeration (CDE for short) of a GMR \( R \), assume that the pairs \( (\bar{t}, R(\bar{t})) \) of \( \text{En} \{ (\bar{t}, R(\bar{t})) \mid \bar{t} \in \text{ supp}(R) \} \) are stored in an array \( A \) (without duplicates). Then \( A \) supports CDE of \( R: \text{enum}(A) \) simply iterates over each element in \( A \), one by one, always outputting the current element. Since array indexing is an \( O(1) \) operation, this gives constant delay. This example shows that CDE of \( Q(db) \) can always be done naively by materializing \( Q(db) \) in an in-memory array. Unfortunately, this requires memory proportional to \( |Q(db)| \) which, depending on \( Q \), can be of size polynomial in |\( db \)|. We hence desire other data structures to represent \( Q(db) \) using less space, while still allowing CDE.

To understand how this can be done, it is instructive to consider a simple binary join \( Q = R(x,y) \Join S(y,z) \) and analyze why traditional join processing algorithms do not yield CDE. Suppose that we evaluate \( Q \) using a simple in-memory hash join with \( R \) as build relation and \( S \) as probe relation. Assume that the corresponding index of \( R \) on \( y \) (i.e., the hash table) has already been computed. Now observe that, when iterating over \( S \) to probe the index, we may have to visit an unbounded number of \( S \)-tuples that do not join with any \( R \)-tuple. Consequently, the delay between consecutive output tuples may be as large as |\( S \)|, which is not constant. A similar analysis shows that other join algorithms, such as the sort-merge join, do not yield enumeration with constant delay.
How can we obtain CDE for $R(x, y) \bowtie S(y, z)$? Intuitively speaking, if we can ensure to only iterate over those $S$-tuples that have matching $R$-tuples, we trivially obtain constant delay since then every probe will yield a new output tuple. As such, the key is to first compute $Y = \pi_y(R) \bowtie \pi_y(S)$ and index both $R$ and $S$ on $y$. We can then iterate over the elements of $Y$, probing both $R$ and $S$ in each iteration to generate the output with constant delay. In the presence of updates, this means that we only need to maintain $Y$, as well as the indexes on $R$ and $S$—all of which are of size linear in $db$ and can be maintained efficiently.

3.2 The Algorithm

To extend the intuition of Section 3.1 from a binary join to general CQs that feature both multway equi-joins and projections, we need to maintain for all the relations that are used as probe relations in a join, the set of tuples that will match the corresponding build relation(s). Of course, we also need to decide in what order we will join the relations, since this determines the auxiliary sets of tuples (like $Y$ above) that we need to maintain. For DYN, this query plan is specified by means of a pair $(T, N)$ called a GJT pair.

GJT pairs. To simplify notation, we denote the set of all variables (resp. atoms, resp. predicates) that occur in an object $X$ (such as a query) by $\text{var}(X)$ (resp. $\text{at}(X)$, resp. $\text{pred}(X)$). In particular, if $X$ is itself a set of variables, then $\text{var}(X) = X$. We extend this notion uniformly to labeled trees. E.g., if $n$ is a node in tree $T$, then $\text{var}(n)$ denotes the set of variables occurring in the label of $n$, and similarly for edges and trees themselves. If $T$ is clear from the context, we omit subscripts from our notation.

Definition 3.1. A GJT pair is a tuple $(T, N)$ with $T$ a generalized join tree and $N$ a sibling-closed connex subset of $T$. A generalized join tree (GJT) is a node-labeled directed tree $T = (V, E)$ such that:
- $T$ is binary: every node has at most two children.
- Every leaf is labeled by an atom.
- Every interior node $n$ is labeled by a hyperedge and has at least one child $c$ such that $\text{var}(n) \subseteq \text{var}(c)$. Such a child is called a guard of $n$.
- Whenever the same variable $x$ occurs in the label of the two nodes $m$ and $n$ of $T$, then $x$ occurs in the label of each node on the unique path linking $m$ and $n$.

A simple GJT is a GJT where $\text{var}(n) \subseteq \text{var}(c)$ for every node $n$ with child $c$, i.e., a GJT where every child is a guard of its parent. A connex subset of $T$ is a set $N \subseteq V$ that includes the root of $T$ such that the subgraph of $T$ induced by $N$ is a tree. $N$ is sibling-closed if for every node $n \in N$ with a sibling $m$ in $T$, $m$ is also in $N$. The frontier of a connex set $N$ is the subset $F \subseteq N$ consisting of those nodes in $N$ that are leaves in the subtree of $T$ induced by $N$.

A GJT pair $(T_1, N_1)$ is simple, but $T_2$ is not since $t(x, u)$ is not a guard of $\{x, y\}$. The set $N_1 = \{\{x\}, \{x, y\}, t(x, u)\}$, highlighted in gray, is a sibling-closed connex subset of $T_1$, and its frontier is $\{\{x, y\}, t(x, u)\}$.

Definition 3.2. Let $(T, N)$ be a GJT pair and assume that \{
\{\pi_1(T)\}, \ldots, \pi_n(T)\}\) is the multiset of atoms occurring as labels in the leaves of $T$. Then the query associated to $T$ is the full join $Q[T] = \pi_1(T) \bowtie \cdots \bowtie \pi_n(T)$ and the query associated to $(T, N)$ is the CQ $Q[T, N] = \pi_{\text{var}(N)}(Q[T])$.

Figure 3: Two example GJT pairs.

Figure 4: $(T_1, N_1)$-representation for the database $db$ specified by the GMRs depicted at the leaves.

The data structure. Following the intuition of Section 3.1, a GJT pair $(T, N)$ acts as query plan by which DYN processes $Q[T, N]$ dynamically. In particular, the GJT $T$ specifies the data structure to be maintained and drives the processing of updates, while the connex set $N$ drives the enumeration of query results. The data structure itself is defined next.

Definition 3.3. Let $(T, N)$ be a GJT pair and let $db$ be a database over $\text{at}(Q)$. The T-reduct (or semi-join reduction) of $db$ is a collection $\rho$ of GMRs, one GMR $\rho_n$ for each node $n \in T$, defined inductively as follows:
- if $n = r(x)$ is an atom, then $\rho_n = db_{r(x)}$.
- if $n$ has a single child $c$, then $\rho_n = \pi_{\text{var}(n)}(\rho_c)$.
- otherwise, $n$ has two children $c_1$ and $c_2$. In this case we have $\rho_n = \pi_{\text{var}(n)}(\rho_{c_1} \bowtie \rho_{c_2})$. Note that, because $n$ has a guard child, this is actually a semijoin.

A T-reduct needs to be augmented by suitable index structures to be used for both enumeration and maintenance under updates. Concretely for each node $n$ with parent $p$ in $T$, the following indexes are created:
- If $n$ belongs to $N$, then we store an index $P_n$ on $\text{var}(p) \bowtie \text{var}(n)$, called the parent index of $n$.
- If $n$ is a node with a sibling $m$, then we store an index $S_n$ of $\text{var}(p) \bowtie \text{var}(m)$, called the sibling index of $n$.

The T-reduct $\rho$ together with the collection of indexes is called a $(T, N)$-representation for $db$, or $(T, N)$-rep for short.

Figure 4 depicts an example $(T_1, N_1)$-representation $\rho$ for the database $db$ composed of the GMRs shown at the leaves of the tree. It is important to observe that the size of this representation for a database $db$ can be at most linear in the size of $db$. The reason is that each interior node only does projections or semijoins. Therefore, as illustrated in Figure 4, for each node $n$ there is some descendant atom
Algorithm 1 Dyn: Dynamic Yannakakis

1: function enumT,N(ρ)
2:   for each \( \vec{t} \in ρ_{root(T)} \) do enumT,N(root(T), \( \vec{t} \), ρ)
3: end for
4: function enumT,N(n, \( \vec{t} \), ρ)
5:   if \( n \) is in the frontier of \( N \) then yield (\( \vec{t} \), \( ρ_n(\vec{t}) \))
6:   else if \( n \) has one child \( c \) then
7:     \( \vec{s} \in ρ \times \vec{t} \) do enumT,N(c, \( \vec{s} \), ρ)
8:   else if \( n \) has two children \( c_1 \) and \( c_2 \) then
9:     for each \( t \in ρ_{c_1} \times \vec{t} \) do enumT,N(c_1, \( t \), ρ)
10:    for each \( t \in ρ_{c_2} \times \vec{t} \) do enumT,N(c_2, \( t \), ρ)
11: end for
12: end if
13: end function

14: procedure update(ρ, u)
15:   for each \( n \in leaves(T) \) do yield (\( \sigma \)) do
16:     \( Δ_n ← u_{\sigma}(\vec{t}) \)
17:   for each \( n \in nodes(T) \setminus leaves(T) \) do
18:     \( Δ_n ← emptyGMRI\over\varsigma(n) \)
19:   for each \( n \in nodes(T) \), traversed bottom-up do
20:     \( ρ_n \leftarrow Δ_n \)
21:     if \( n \) has a parent \( p \) and a sibling \( m \) then
22:       \( β \leftarrow π_{\varsigma}(\vec{t} \setminus \vec{m}) \cup \vec{m} \)
23:     end if
24:     if \( n \) has parent \( p \) then
25:       \( Δ_{p,t} ← π_{\varsigma}(\vec{t} \setminus \vec{m}) \)
26:     end if
27: end procedure

α (possibly \( n \) itself) such that \( supp(ρ_n) ⊆ supp(π_{\varsigma}(\vec{t})db_n) \). Consequently, the indexes are also of size linear in \( db \).

Given these definitions, the enumeration and maintenance algorithms that form the Dynamic Yannakakis algorithm are shown in Algorithm 1. They operate as follows.

Enumeration. To enumerate from a (\( T, N \))-rep we iterate over the reductions \( ρ_n \) with \( n \in N \) in a nested fashion, starting at the root and proceeding top-down. When \( n \) is the root, we iterate over all tuples in \( ρ_n \). For each such tuple \( \vec{t} \), we iterate only over the tuples in the children \( c \) of \( n \) that are compatible with \( \vec{t} \) (i.e., tuples in \( ρ \) that join with \( \vec{t} \)). Note that such tuples can be enumerated efficiently thanks to the index \( P \). This procedure continues until we reach nodes in the frontier of \( N \) at which time the output tuple can be constructed.

The pseudocode is given by the routine \( enum \) in Algorithm 1, where the tuples that are compatible with \( \vec{t} \) are computed by \( ρ \times \vec{t} \).

Update processing. To maintain a (\( T, N \))-rep under update \( u \) it suffices to traverse the nodes of \( T \) in a bottom-up fashion. At each node \( n \) we have to compute the update \( Δ_n \) to apply to \( ρ_n \) and its associated indexes. For leaf nodes, this update is given by the update \( a \) itself. For interior nodes, \( Δ_n \) can be computed from the update and the original reduct of its children. Algorithm 1 gives the pseudocode. Here, line 21 is then implemented by means of a straightforward hash-join (using the sibling index \( S_m \) on \( ρ_m \)). As a side effect of modifying \( p \) the associated indexes are also updated (not shown).

Theorem 3.4. Let (\( T, N \)) be a fixed GJT p air. Given a (\( T, N \))-rep of \( db \) with \( T \)-reduct \( ρ \), \( enumT,N(ρ) \) enumerates \( Q[T,N](db) \) with constant delay. Moreover, \( update(ρ, u) \) updates the (\( T, N \))-rep from a (\( T, N \))-rep of \( db \) to a (\( T, N \))-rep of \( db + u \) in time \( O(|db| + |u|) \). If \( T \) is simple, then the update time is \( O(|u|) \), hence independent of \( |db| \).

Note that \( update(ρ, u) \) can be used to build a (\( T, N \))-rep of \( db \) in time \( O(|db|) \): start from an empty (\( T, N \))-rep (which represents the empty database) and then call \( update(ρ, u) \) with \( u = db \). This hence shows that Dyn can be used to enumerate \( Q[T,N](db) \) with constant delay after linear time preprocessing.

We also note that if \( \vec{t} \) is a tuple over \( var(M) \) for some connex subset \( M ⊆ N \) of \( T \), then checking whether \( \vec{t} \in π_{\varsigma}(\vec{t})Q[T,N](db) \) can be done in constant time: it suffices to check that \( \vec{t}[\varsigma(M)] \in ρ_n \) for every \( m \in M \) and return true if and only if this is the case. Since \( T \) and \( N \) are fixed, the size of \( M \) is bounded and these are a constant number of checks, all of which run in constant time.

Discussion. Dyn heavily relies on having a GJT pair (\( T, N \)) to process queries. If, for a CQ \( Q \) there exists some GJT pair (\( T, N \)) such that \( Q = Q[T,N] \) then \( Q \) is said to be free-connex acyclic, and (\( T, N \)) is called a GJT pair for \( Q \). If \( Q \) is acyclic if full(\( Q) = Q[T,N] \) for some \( T \).

Not all CQs are (free-connex acyclic). For instance, the triangle query \( r(x, y) \land s(y, z) \land t(x, z) \) is the prototypical example of a non-acyclic query. Furthermore, \( π_{x,z}(r(x, y) \land s(y, z)) \) is acyclic but not free-connex acyclic. Since every free-connex acyclic CQ is acyclic, this example shows that free-connex acyclic CQs form a strict subclass of the acyclic CQs. Recent analysis of query logs show that free-connex acyclic queries occur very frequently in practice [5]. We refer readers interested in algorithms for computing GJT pairs for CQs [14].

One may wonder whether algorithms with the same properties as Dyn can be obtained for CQs that are not free-connex acyclic. It is known that it is not possible for the class of all acyclic CQs, unless multiplication of two \( n \times n \) binary matrices can be computed in \( O(n^2) \) time [3]. Using further complexity-theoretic assumptions, it is possible to show that this is also not possible for the class of all CQs [6].

It is also known [4] that, unless the Online Matrix-Vector Multiplication conjecture [11] is false, the class of queries that allow both (1) constant-delay enumeration of query results and (2) update processing time \( O(|u|) \) for every update \( u \), is exactly the class of so-called \( q \)-hierarchical queries. While we forego a formal definition of this class, we show in [12] that a CQ \( Q \) is \( q \)-hierarchical if, and only if there exists a GJT pair (\( T, N \)) for \( Q \) such that \( T \) is simple. Since Dyn has update time \( O(|u|) \) for exactly these queries, Dyn hence meets the theoretical lower bound.

For readers familiar with the Yannakakis algorithm [23] it may not be obvious from the description above why Dyn can be claimed to be a dynamic version of Yannakakis. We refer to [12] for a discussion.

4 Dealing with \( θ \)-joins

To extend Dyn to also process \( θ \)-joins, it is instructive to consider the GJT \( T = (R(x, y) \land S(y, z) \mid x < z) \) where the \( θ \)-join is an inequality-join. To obtain CDE for \( Q \), assume that we have already computed \( Y = π_{x,y}(\sigma_{x<z}(R(x, y) \land S(y, z))) \) and that, moreover, we have a more powerful index structure \( I \) that allows, for any tuple \( \{x, y\} \)-tuple \( \vec{t} \) over, to enumerate \( \sigma_{x<z}(S(y, z) \times \vec{t}) \) with constant delay. We can then obviously enumerate \( Q \) with constant delay by iterating

There exists many equivalent definitions of when a join query is acyclic, and consequently also of when a CQ with projections is free-connex acyclic. See [12, 14] for a discussion of why the new definition that we give here is equivalent to the existing ones.
over the elements of \( \vec{t} \in Y \), and for each such \( \vec{t} \), probe \( I \) to produce the tuples \( \vec{s} \in \sigma_{x<z}(S(y,z) \triangleright \vec{t}) \), outputting each \( \vec{s} \cup \vec{t} \). Since \( \sigma_{x<z}(S(y,z) \triangleright \vec{t}) \) allows CDE, the entire procedure is CDE. The key question then, is how we can build this more powerful index structure \( I \). The solution is to build a normal (hash-based) index \( J \) of \( S \) on \( y \) but use sorting to store the index in such a way that for every \( \vec{t} \) over \( y \) this index returns a pointer to \( S \times \vec{t} \) for which tuples are enumerated in descending order on \( z \). This now supports CDE of \( \sigma_{x<z}(S(y,z) \triangleright \vec{t}) \), for every \( \vec{t} \) over \( x,y \): use \( J \) to enumerate \( R \times \vec{t} \) with constant delay and in decreasing order on \( z \). Yield the current tuple \( \vec{s} \) that is being enumerated in this fashion, provided that \( I(x) < z(\vec{s}) \). As soon as \( I(x) \geq z(\vec{s}) \) we know that all subsequent \( \vec{s} \) will fail the inequality, and we can hence terminate. This example forms the basic intuition how we can extend DYN to deal with GGCQs with inequality joins. We next sketch the generalization to arbitrary \( \theta \)-joins, and refer to [13] for detailed exposition.

First, in the presence of \( \theta \)-joins, a GJT pair is defined exactly as in Definition 3.1 except that now additionally every edge \( p \rightarrow c \) from parent \( p \) to child \( c \) is labeled by a set \( \text{pred}(p \rightarrow c) \) of predicates. It is required that every predicate \( \theta(\vec{t}) \) in this set satisfies \( \forall var(p) \cup \forall var(c) \). The query \( Q[T,N] \) associated to \((T,N) \) then becomes \( \pi_{\theta(\vec{t})} (\bigcup_{\vec{t} \in \text{pred}(T)} Q[T]) \). Here, \( \text{pred}(T) \) are all the predicates occurring on edges in \( T \).

Next, \((T,N)\)-reps are extended to account for predicates. Concretely, the inductive definition of \( \rho_n \) in the \( T \)-reduce becomes:

\[-\text{if } n = \tau(x) \text{ is an atom, then } \rho_n = db_r(x)\]
\[-\text{if } n \text{ has a single child } c, \text{ then } \rho_n = \pi_{\forall var(n)}(\text{pred}(n \rightarrow c)) \rho_c\]
\[-\text{otherwise, } n \text{ has two children } c_1 \text{ and } c_2. \text{ In this case we set } \rho_n = \pi_{\forall var(n)}(\text{pred}(n \rightarrow c_1) \rho_{c_1} \land \text{pred}(n \rightarrow c_2) \rho_{c_2}).\]

Here \( \text{pred}(n) \) denotes the set of all predicates on the edges from \( n \) to its children in \( T \). The indexes that we need to maintain are modified as follows: \( P_n \) should now allow CDE of \( \sigma_{\text{pred}(n \rightarrow m)}(\rho_m \triangleright \vec{t}) \), for every \( \vec{t} \) over \( \forall var(p) \), where \( p \) is \( n \)'s parent. \( S_n \) should allow CDE of \( \sigma_{\text{pred}(n \rightarrow p)}(\rho_p \triangleright \vec{t}) \) for every tuple \( \vec{t} \) over \( \forall var(m) \) where \( m \) is the sibling of \( n \). The exact design of these indexes of course depends on the semantics of the predicates included in \( T \); for inequality predicates we have sketched above how they work. The generalization of DYN works as long as we have these indexes.

Finally, Algorithm 1 is modified so that in Lines 6, 8 and 9 we iterate over \( \sigma_{\text{pred}(n \rightarrow m)}(\rho_m \triangleright \vec{t}) \) resp. \( \sigma_{\text{pred}(n \rightarrow c_1)}(\rho_{c_1} \triangleright \vec{t}) \) and \( \sigma_{\text{pred}(n \rightarrow c_2)}(\rho_{c_2} \triangleright \vec{t}) \). Lines 19, 21, 23 are modified to compute the \( \Delta \) GMRs under the now-modified definition of \((T,N)\)-rep. We refer to the general version of DYN with arbitrary \( \theta \)-joins as GDYN, and the version where all \( \theta \)-joins are inequalities as IEDYN.

**Theorem 4.1.** Let \((T,N)\) be a fixed GJT pair. Given a \((T,N)\)-rep of \( db \), both GDYN and IEDYN correctly enumerate \( Q[T,N](db) \) and update the \((T,N)\)-rep to a \( db \) under update \( u \). In the case that all predicates in \( \theta \)-joins are inequalities, IEDYN has the following complexity. If there is at most one inequality on each edge in \( T \), then IEDYN enumerates with constant delay and has \( O(M \cdot \log(M)) \) update time where \( M = |db| + |u| \). If \( T \) has some edge that contains multiple inequalities, the delay is \( O(\log(|db|)) \) and the update time is \( O(M^2 \cdot \log(M)) \).\(^2\)

\(^2\)In [13] there was an incorrect claim: we stated that updates could be processed in time \( O(M \cdot \log(M)) \) in this last case. We then found a bug in our algorithm and we currently do not know if this bound can be achieved. See [14] for a proof that IEDYN runs in the bounds claimed here.

## 5 Experimental Evaluation

We have implemented (IE)\( DYN \) as a query compiler that generates executable code in the Scala programming language. The generated code instantiates a \((T,N)\)-rep for a query \( Q \) and defines trigger functions that are used for maintaining the \((T,N)\)-rep under updates.

Our implementation supports two modes of operation: push-based and pull-based. In both modes, the system maintains the \((T,N)\)-rep under updates. In the push-based mode the system generates, on its output stream, the delta result \( \Delta Q(db, u) := Q(db + u) - Q(db) \) after each single-tuple update \( u \). To do so, it uses a modified version of enumeration that we call delta enumeration. Similarly to how enum enumerates \( Q(db) \), delta enumeration enumerates \( \Delta Q(db, u) \) with constant delay (if \( Q \) has at most one inequality per pair of atoms) resp. logarithmic delay otherwise. To do so, it uses both (1) the \((T,N)\)-reduce GMRs \( \rho_n \) and (2) the delta GMRs \( \Delta \rho_n \) that are computed by \text{update} when processing \( u \). In this case, however, one also needs to index the \( \Delta \rho_n \) similarly to \( \rho_n \). In the pull-based mode, in contrast, the system only maintains the \((T,N)\)-rep under updates but does not generate any output stream. Nevertheless, at any time a user can call enum to obtain the current output.

It should be noted that our implementation also supports the processing of general acyclic GGCQs that are not necessarily free-connex. This is done using the following simple strategy. Let \( Q \) be acyclic but not free-connex. First, compute a free-connex acyclic approximation \( Q_F \) of \( Q \). \( Q_F \) can always be obtained from \( Q \) by extending the set of output variables of \( Q \). In the worst case, we need to add all variables, and \( Q_F \) becomes the full join underlying \( Q \). Then, use (IE)\( DYN \) to maintain a \((T,N)\)-rep for \( Q_F \). When operating in push-based mode, for each update \( u \), we use the \((T,N)\)-rep to delta-enumerate \( \Delta Q_F(db, u) \) and project each resulting tuple to materialize \( \Delta Q(db, u) \) in an array. Subsequently, we copy this array to the output. Note that the materialization of \( \Delta Q(db, u) \) here is necessary since the delta enumeration can produce duplicate tuples after projection. When operating in pull-based mode, we materialize \( Q(db) \) in an array, and use delta enumeration of \( Q_F \) to maintain the array under updates. Of course, under this strategy, we require \( \Omega(|Q(db)|) \) space in the worst case, just like (J)IVM would, but we avoid the (partial) materialization of delta queries. Note the distinction between the two modes: in push-based mode \( \Delta Q(db, u) \) is materialized (and discarded once the output is generated), while in pull-based mode \( Q(db) \) is materialized upon requests. Finally, our implementation also supports common aggregates like \text{SUM} and \text{AVG}, see [12] for more information.

### 5.1 Conjunctive Queries

We evaluate a subset of queries available in the industry-standard benchmarks TPC-H and TPC-DS. In particular, we evaluate those queries involving only equijoins, whose FROM-WHERE clauses are acyclic. Queries are divided into acyclic full-join queries (called FQs) and acyclic aggregate queries. Acyclic full join queries are generated by taking the FROM clause of the corresponding queries on the benchmarks. We omit the ORDER BY and LIMIT clauses, we replaced the
left-outer join in TPC-H query Q13 by an equijoin, and modified TPC-H Q16 to remove an inequality. See [12] for the full query specification.

Our workload consist of a stream of updates, where each update consists of a single-tuple insertion. The streams were generated using the TPC-H and TPC-DS data generators. The number of tuples in each stream is depicted in Table 1.

We compare IEDYN with DBToaster [16] using memory footprint and update processing time as comparison metrics. DBToaster is a state-of-the-art implementation of HIVM. It operates in pull-based mode, and is optimized for aggregations over equi-joins. DBToaster has been extensively tested for such queries and has proven to be more efficient than a commercial database management system, a commercial stream processing system and an IVM implementation [16]. It is therefore an interesting implementation to compare to. DBToaster compiles given SQL statements into executable trigger programs in different programming languages. We compare against those generated in Scala from the DBToaster Release 2.2.³

³https://dbtoaster.github.io/

Figure 5 depicts the resources used by DYN as a percentage of the resources consumed by DBToaster (lower is better). From Figure 5 we see that for full join queries (FQ1-FQ5), DYN outperforms DBToaster by close to one order of magnitude in both memory consumption and processing time, illustrating the effectiveness of maintaining (T,N)-reps rather than the query results themselves, especially when these results are large. For aggregate queries, Figure 5 shows that DYN can significantly improve the memory consumption of HIVM while improving processing time—up to two orders of magnitude for TPC-H Q13³ and TPC-DS Q7. See [12] for an in-depth discussion.

While the T-reps maintained by IEDYN feature constant delay enumeration, this theoretical notion hides a constant factor that could decrease performance in practice when compared to full materialization. Experiments detailed in [12] show that this not the case: DYN’s enumeration time is competitive with DBToaster.

### 5.2 Conjunctive Queries with Inequalities

To gauge the effectiveness of IEDYN on GCQs that feature inequality joins, we evaluate the acyclic queries listed in Table 2 on synthetically-generated streams of single-tuple insertion updates. The sizes of the update streams are intentionally kept low, since they generate huge output sizes (cf. Table 2).

Here we only compare IEDYN with Esper ⁴ but refer to [13] for a more detailed comparison against other state of the art

⁴http://www.espertech.com/esper/esper-downloads/

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Table 1: CQ benchmark stream sizes.

<table>
<thead>
<tr>
<th>Benchmark</th>
<th>Query</th>
<th># of tuples</th>
</tr>
</thead>
<tbody>
<tr>
<td>TPC-H</td>
<td>FQ1</td>
<td>2,833,827</td>
</tr>
<tr>
<td></td>
<td>FQ2</td>
<td>2,031,163</td>
</tr>
<tr>
<td></td>
<td>FQ4</td>
<td>2,820,494</td>
</tr>
<tr>
<td></td>
<td>FQ5</td>
<td>2,704,494</td>
</tr>
<tr>
<td></td>
<td>Q1</td>
<td>7,999,406</td>
</tr>
<tr>
<td></td>
<td>Q3</td>
<td>10,199,406</td>
</tr>
<tr>
<td></td>
<td>Q4</td>
<td>9,999,406</td>
</tr>
<tr>
<td></td>
<td>Q6</td>
<td>9,999,406</td>
</tr>
<tr>
<td></td>
<td>Q9</td>
<td>11,396,069</td>
</tr>
<tr>
<td></td>
<td>Q12</td>
<td>9,999,406</td>
</tr>
<tr>
<td></td>
<td>Q13</td>
<td>2,200,000</td>
</tr>
<tr>
<td></td>
<td>Q16</td>
<td>1,333,339</td>
</tr>
<tr>
<td></td>
<td>Q18</td>
<td>10,199,406</td>
</tr>
<tr>
<td>TPC-DS</td>
<td>FQ5</td>
<td>10,669,570</td>
</tr>
<tr>
<td></td>
<td>Q3</td>
<td>11,635,973</td>
</tr>
<tr>
<td></td>
<td>Q7</td>
<td>13,591,299</td>
</tr>
<tr>
<td></td>
<td>Q19</td>
<td>11,987,115</td>
</tr>
<tr>
<td></td>
<td>Q22</td>
<td>30,138,621</td>
</tr>
</tbody>
</table>

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https://doi.org/10.1145/3371316.3371325

Provided by Sächsische Landesbibliothek, Staats- und Universitätsbibliothek Dresden
systems. Esper is a complex event processing engine with a relational model based on Stanford STREAM [2]. It operates in push-based mode. We use the Java-based open source implementation4 for our comparisons.

Figure 6 depicts the resources used by IEDYN as a percentage of the resources used by Esper, both operating in push-based mode. IEDYN significantly outperforms Esper on all full join queries (GCQ1–GCQ4). We note that for these queries, even in push-based mode IEDYN can support the enumeration of query results from its data structures at any time while competing push-based systems have no such support. Hence, IEDYN is not only more efficient but also provides more functionality. IEDYN also significantly outperforms Esper on free-connex queries GCQ5 and GCQ6 with more than a threefold improvement in processing time and an order of magnitude improvement in memory usage on Q7. For non-free-connex queries GCQ7 and GCQ8, IEDYN continues to significantly outperform Esper in processing time, showing an order of magnitude improvement in memory usage for GCQ7.

### 6 Summary

Traditional techniques for dynamic query evaluation are based either on materialization (to avoid recomputation of subresults), or on recomputation (to avoid the space overhead of materialization). We have shown that both techniques are suboptimal: instead of materializing subresults, one can use Dynamic Yannakakis to maintain a data structure that is succinct; and yet supports all operations one commonly expects from materialization: enumeration with constant delay as well as fetching single tuples in constant time. Our experiments against state-of-the-art engines in different domains show that this can improve performance by orders of magnitude.

### 7 References