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Applications of Random Sampling to On-line Algorithms in Computational Geometry

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Abstract

This paper presents a general framework for the design and randomized analysis of geometric algorithms. These algorithms are on-line and the framework provides general bounds for their expected space and time complexities when averaging over all permutations of the input data. The method is general and can be applied to various geometric problems. The power of the technique is illustrated by new efficient on-line algorithms for constructing convex hulls and Voronoï diagrams in any dimension, Voronoï diagrams of line segments in the plane, arrangements of curves in the plane and others.

1 Introduction

Randomized incremental construction is a new paradigm in Computational Geometry that has been successfully applied to a variety of problems [CS89, Mul89, MMO91]. These algorithms are rather simple, easy to code and efficient in practice. Moreover, they do not require that the input data satisfy some probabilistic distribution but only that they are inserted in random order.

Clarkson and Shor [CS89] have given a general framework for the design and the analysis of such randomized incremental constructions. In this framework, geometrical problems are stated in terms of objects, regions and conflicts between objects and regions. The input data form a set of objects, the regions are defined by small subsets of objects and each region is in conflict with a (possibly empty) subset of the input objects. Precise definitions of objects, regions and conflicts of course depend of the particular problem at hand. The output of the problem is assumed to be closely related to the set of regions defined by the input objects which are not in conflict with those objects. The algorithm incrementally constructs the set of regions defined by a current subset of the input objects which are not in conflict with that subset and maintains in an additional data structure, called the conflict graph, the conflicts between the regions of the current construction and the objects not yet introduced. When a new object is to be inserted, the regions in conflict with that object are replaced by new regions and the conflict graph is updated. This general framework has been shown to yield efficient algorithms to such problems as computing the convex hull or the Voronoï diagram of a set of points or computing the intersections of a set of line segments.

Because of the conflict graph, these algorithms are static in nature. A few attempts have been proposed to avoid the conflict graph and thus to obtain on-line algorithms in some specific applications [BDT, BT86, GKS92]. The present paper extends these results and provides a general framework for on-line algorithms with good expected behaviour when the objects are assumed to be introduced in random order. The basic idea, first introduced in [BT86], is the following. Instead of replacing old regions by new ones, we store, into a data structure called the Influence DAG, all the regions that appear at some stage of the construction as defined by the current subset of objects and without conflict with that subset. This structure is constructed on-line and designed to efficiently retrieve the conflicts between any new object to be inserted and the current set of regions.

Our main result bounds the cost of inserting a new object into a set of already inserted objects. Our results hold when averaging over all possible permutations of the set of inserted objects.

In the second part of the paper, we show that the major applications that can be solved using the Clarkson and Shor approach can also be solved using our framework. For all these applications, we obtain simple on-line algorithms with

the same complexity as their static counterparts.

More specifically, the convex hull of a set of n points can be computed with logarithmic cost per added point in dimensions 2 and 3 and in time $O(n^{\lfloor \frac{d}{2} \rfloor - 1})$ in dimension $d > 3$. By the well known correspondence between Delaunay triangulations (and Voronoï diagrams) in dimension d and convex hulls in dimension $d + 1$, we immediately deduce algorithms for constructing those structures. A more direct construction can also be derived and further extended to compute the Voronoï diagram of a set of line segments in the plane at the same cost. All these complexities are worst-case optimal for randomized algorithms.

Another application consists in computing an arrangement of planar curves (of bounded degree) in time $O(\log n + a/n)$ per insertion using $O(n + a)$ space, where a is the size of the output. This algorithm can also be used to locate a point in a planar subdivision in $O(\log n)$ query time after a preprocessing that requires $O(n \log n)$ time and $O(n)$ space.

This list is not exhaustive and several other applications will be mentioned.

2 The general framework

2.1 Definitions and notations

Following Clarkson and Shor [CS89], we formulate the geometric problems considered in this paper in terms of *objects* and *regions*.

The *objects* are members of a universe \mathcal{O} and are the input data of the problem. Typically, objects are points, segments, lines, planes or hyperplanes in a usual d -dimensional euclidean space E^d .

The *regions* are defined by subsets of the universe \mathcal{O} of cardinality less than a constant b . This assumes that there is a relation between the subsets of \mathcal{O} of cardinality less than b and the regions; a region F is then said to be defined by a subset \mathcal{X} of objects if \mathcal{X} is in relation with F . Moreover, as in [CS89], we define *conflicts* between objects and regions. The subset of objects of \mathcal{O} which are in conflict with a region is called the *influence range* of the region. Conflicts are defined in such a way that a region F defined by the subset \mathcal{X} is not in conflict with the objects of \mathcal{X} .

Of course, the notion of conflict or equivalently the influence range of a region F has to be made precise for each specific problem. Still following Clarkson, we assume that, within this framework, the required geometric construction can be formulated as the following problem : given a finite subset \mathcal{S} of the object universe \mathcal{O} , find all the regions that are defined by objects in \mathcal{S} and that have no object of \mathcal{S} in their influence range.

For a finite set of objects \mathcal{S} we denote by $\mathcal{F}(\mathcal{S})$ the set of regions defined by the objects in \mathcal{S} . We call the number of objects of \mathcal{S} that belong to the influence

range of F width of a region F with respect to \mathcal{S} . Let $\mathcal{F}_j(\mathcal{S})$ be the subset of $\mathcal{F}(\mathcal{S})$ consisting of the regions that have width j with respect to \mathcal{S} .

2.2 The I-DAG (Influence Directed Acyclic Graph)

The different algorithms presented in this paper are incremental and introduce objects one by one. Let \mathcal{S} be the set of objects which have already been introduced. At a given stage, the incremental algorithm inserts a new object in \mathcal{S} and updates the set $\mathcal{F}_0(\mathcal{S})$ of regions of zero width defined by \mathcal{S} . This is performed through the maintenance of a dynamic structure called the Influence DAG (I-DAG for short) described below.

The I-DAG is a rooted directed acyclic graph whose nodes are associated with regions that, at some stage of the algorithm, have appeared as regions of zero width defined by the set of objects that have been introduced at that stage. Although the I-DAG is not strictly speaking a tree, we speak of leaves, fathers, sons etc. in the obvious way. The nodes of the I-DAG associated with regions of current zero width with respect to \mathcal{S} are marked. When a new object O is added to \mathcal{S} , one new node is created for each region in $\mathcal{F}_0(\mathcal{S} \cup \{O\})$ that is not a region of $\mathcal{F}_0(\mathcal{S})$.

The already existing nodes are never deleted from the I-DAG but possibly unmarked. The new nodes are linked by edges to other nodes of the I-DAG. These edges are constructed in such a way that a new object O can be efficiently located in the structure ; locating O means here to find all the nodes whose associated regions are in conflict with O . When a new node corresponding to a region F of $\mathcal{F}_0(\mathcal{S} \cup \{O\})$ is added to the I-DAG, we put edges between already existing nodes and the new nodes so that the influence range of F is included in the union of the ranges of its parents.

The I-DAG structure is characterized by the two following fundamental properties :

Property 1 At each stage of the incremental process, the regions of zero width ($\mathcal{F}_0(\mathcal{S})$) are leaves of the I-DAG.

Property 2 The influence range of the region associated with a node is included in the union of the ranges of its parents.

The construction of the I-DAG can be sketched as follows:

- We initialize \mathcal{S} with the b first objects. A node of the I-DAG is created for each region of $\mathcal{F}_0(\mathcal{S})$ and made son of the root of the I-DAG. The influence range associated with the root is the whole objects universe \mathcal{O} .
- At each subsequent step, a new object O is added to \mathcal{S} and the I-DAG is updated. The two following substeps are performed:

location substep. This substep finds all the nodes of the I-DAG whose regions have zero width and are in conflict with O . This is done by traversing every path from the root of the I-DAG down to the first node which is not in conflict with object O .

creation substep. From the information collected during the location substep, the creation substep creates a new node for each region in the set $\mathcal{F}_0(\mathcal{S} \cup \{O\}) \setminus \mathcal{F}_0(\mathcal{S})$ and links the new nodes to already existing nodes in the structure so that Properties 1 and 2 still hold. The details of this substep depend on each particular application.

2.3 Randomized analysis of the I-DAG

This subsection aims to provide a randomized analysis of the space and time required to build the I-DAG structure. Randomization concerns here only the order in which the inserted objects are introduced in the structure. Thus, if the current set of objects is a set S of cardinality n , our results are expected values that correspond to averaging over the $n!$ possible permutations of the inserted objects, each equally likely to occur.

We first prove some probabilistic results that are purely combinatorial. These results allow us to prove the main results of this paper, stated as Theorems 2.4 and 2.9 below, that give quite general upper bounds on the size and the update time of the I-DAG as functions of the expected size of the output for a sample of the input.

For the sake of clarity, we will make some hypotheses that simplify somewhat the analysis. These conditions are fulfilled by a large class of geometric problems and allow one to express the results in a simple way. However these hypothesis are not really necessary and will be removed in Section 2.4.

Probabilistic Lemma

We first introduce some additional notation and definitions. In some geometric situations (e.g. when computing an arrangement of line segments) the regions are not all defined by the same number of objects. We adapt our notation accordingly : $\mathcal{F}_j^i(\mathcal{S})$ will denote the subset of regions of $\mathcal{F}(\mathcal{S})$ defined by i objects of \mathcal{S} ($i \leq b$) and having width j , and $\mathcal{F}_{\leq j}^i(\mathcal{S})$ will denote the subset of regions defined by i objects and having width at most j .

A r -random sample of a finite set of objects \mathcal{S} of cardinality n , is a subset of \mathcal{S} of r objects chosen at random, i.e. with probability $\frac{1}{\binom{n}{r}}$.

Let $f_0(r, \mathcal{S})$ be the expected size $E(|\mathcal{F}_0(\mathcal{R})|)$ of $\mathcal{F}_0(\mathcal{R})$ for random samples \mathcal{R} of cardinality r .

The first Lemma, due to Clarkson and Shor [CS89], bounds the number $|\mathcal{F}_{\leq j}^i(\mathcal{S})|$ of regions with width at most j defined by \mathcal{S} . We recall it for completeness.

Lemma 2.1

$$|\mathcal{F}_{\leq j}^i(\mathcal{S})| = O\left(j^i f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$$

Proof. The proof uses the technique of random sampling. Let \mathcal{R} be a r -random sample of \mathcal{S} . A region F of $\mathcal{F}_j^i(\mathcal{S})$ has width zero with respect to \mathcal{R} if and only if the i objects defining F belong to \mathcal{R} while the j objects in conflict with F do not. So the probability $\text{Prob}(F \in \mathcal{F}_0(\mathcal{R}))$ that F belongs to $\mathcal{F}_0(\mathcal{R})$ is :

$$\begin{aligned} \text{Prob}(F \in \mathcal{F}_0(\mathcal{R})) &= \frac{\binom{n-i-j}{r-i}}{\binom{n}{r}} \\ &= \frac{(n-i-j)!r!(n-r)!}{(n-r-j)!(r-i)!n!} \\ &= \frac{r \dots (r-i+1) (n-r) \dots (n-r-j+1)}{n \dots (n-i+1) (n-i) \dots (n-i-j+1)} \\ &= \frac{r \dots (r-i+1)}{n \dots (n-i+1)} \left(\frac{(n-r-k+1)}{(n-i-k+1)} \right)^k \text{ for } j \leq k \\ &\geq \frac{r \dots (r-i+1)}{n \dots (n-i+1)} \left(1 - \frac{1}{k} \right)^k \text{ if } r = \lfloor \frac{n}{k} \rfloor, \text{ for } 1 \leq k \\ &\geq \frac{1}{4} \frac{r \dots (r-i+1)}{n \dots (n-i+1)} \text{ for } k \geq 2 \end{aligned}$$

Then we compute the expected size of $\mathcal{F}_0^i(\mathcal{R})$.

$$\begin{aligned} E(|\mathcal{F}_0^i(\mathcal{R})|) &= \sum_{j=0}^{n-i} \sum_{F \in \mathcal{F}_j^i(\mathcal{S})} \text{Prob}(F \in \mathcal{F}_0^i(\mathcal{R})) \\ &\geq \sum_{j=0}^k \sum_{F \in \mathcal{F}_j^i(\mathcal{S})} \text{Prob}(F \in \mathcal{F}_0^i(\mathcal{R})) \\ &\geq \left(\sum_{j=0}^k |\mathcal{F}_j^i(\mathcal{S})| \right) \frac{1}{4} \frac{r \dots (r-i+1)}{n \dots (n-i+1)} \end{aligned}$$

By introducing f_0 , we get :

$$\begin{aligned} |\mathcal{F}_{\leq k}^i(\mathcal{S})| &\leq 4 \frac{n \dots (n-i+1)}{r \dots (r-i+1)} E(|\mathcal{F}_0^i(\mathcal{R})|) \\ &\leq O\left(k^i f_0\left(\lfloor \frac{n}{k} \rfloor, \mathcal{S}\right)\right) \end{aligned}$$

□

The update conditions

We first assume that three update conditions are satisfied. As already mentioned these hypotheses are mainly for clarity and will be removed later.

1. The number of sons of a node of the I-DAG is bounded.
2. Given a region F and an object O , the test to decide whether or not O is in conflict with F can be performed in constant time.
3. If the new object O added to the current set \mathcal{S} is found to be in conflict with k regions of $\mathcal{F}_0(\mathcal{S})$ then the creation substep requires $O(k)$ time.

Expected storage

Lemma 2.2 *If \mathcal{S} has cardinality n , the expected size of the I-DAG of \mathcal{S} is :*

$$O\left(\sum_{j=1}^n \frac{f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})}{j}\right)$$

Proof. The expected number of nodes $\eta(\mathcal{S})$, in the I-DAG of \mathcal{S} can be obtained by summing, for all the regions F of $\mathcal{F}(\mathcal{S})$, the probability that F occurs as a node in the I-DAG, which is $\frac{i!j!}{(i+j)!}$ (the i objects defining F must be inserted before the j objects in F). By Lemma 2.1,

$$\begin{aligned} \eta(\mathcal{S}) &= \sum_{i=1}^b \sum_{j=0}^{n-i} |\mathcal{F}_j^i(\mathcal{S})| \frac{i!j!}{(i+j)!} \\ &= \sum_{i=1}^b \left(|\mathcal{F}_0^i(\mathcal{S})| + \sum_{j=1}^{n-i} (|\mathcal{F}_{\leq j}^i(\mathcal{S})| - |\mathcal{F}_{\leq (j-1)}^i(\mathcal{S})|) \frac{i!j!}{(i+j)!} \right) \\ &= \sum_{i=1}^b \left(\sum_{j=1}^{n-i-1} |\mathcal{F}_{\leq j}^i(\mathcal{S})| i \frac{i!j!}{(i+j+1)!} + |\mathcal{F}_{\leq (n-i)}^i(\mathcal{S})| \frac{i!(n-i)!}{n!} \right) \\ &\leq O\left(\sum_{j=1}^n \frac{f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})}{j}\right) \end{aligned}$$

According to update condition 1, this bound applies also to the size of the I-DAG.
 \square

Expected time

Lemma 2.3 *Under the update conditions, if \mathcal{S} has cardinality n , the expected time for inserting the last object in the I-DAG is*

$$O\left(\frac{1}{n} \sum_{j=1}^n f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$$

Proof. Under update condition 2, the computing time spent to locate of the last inserted object O , is proportional to the total number of nodes of the I-DAG visited when locating O . Due to update condition 1, the number of nodes visited when locating O is at most proportional to the number of nodes of the I-DAG associated with regions in conflict with O . Thus the expected time for locating the last inserted object O is at most proportional to the expected number, $\theta(\mathcal{S})$, of nodes of the I-DAG associated with regions in conflict with O .

Let F be a region of $\mathcal{F}_j^i(\mathcal{S})$. F is a region in conflict with O associated with a node of the I-DAG, if O is one of the j objects in conflict with F and if the i objects defining F have been inserted before the j objects in conflict with F . This occurs with the probability: $\frac{j \cdot i!(j-1)!}{n(i+j-1)!}$. The expected number $\theta(\mathcal{S})$ of nodes visited during the last insertion is then obtained by summing, for all the regions F of $\mathcal{F}(\mathcal{S})$, the above probability. Using Lemma 2.1, this yields :

$$\theta(\mathcal{S}) = \sum_{i=1}^b \sum_{j=1}^{n-i} |\mathcal{F}_j^i(\mathcal{S})| \frac{i!j!}{n(i+j-1)!} = O\left(\frac{1}{n} \sum_{j=1}^n f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right),$$

from a calculation similar to the proof of Lemma 2.2 \square

Due to update condition 3, the computing time of the last creation substep is also dominated by a term proportional to the number of nodes of the I-DAG associated with a region in conflict with O and admits the same expected upper bound as $\theta(\mathcal{S})$.

Main theorem

Lemmas 2.2 and 2.3 prove our main result :

Theorem 2.4 *If the set of already inserted objects \mathcal{S} has cardinality n and if the update conditions are fulfilled, the I-DAG of \mathcal{S} requires $O\left(\sum_{j=1}^n \frac{f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})}{j}\right)$*

expected memory space. The insertion of a new object can be done in $O\left(\frac{1}{n} \sum_{j=1}^n f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$ expected update time.

Corollary 2.5 Under the update conditions, the total expected time to build a I-DAG for a set \mathcal{S} of n object is : $O\left(\sum_{j=1}^n f_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$

Proof. Notice that in that corollary, \mathcal{S} is no longer the current subset of inserted objects, but the final set of objects. From Theorem 2.4, we know that, if a subset \mathcal{R} of \mathcal{S} with cardinality r has been inserted at a given time, the expected time to insert the last object of \mathcal{R} is :

$$O\left(\sum_{j=1}^r \frac{1}{r} f_0(\lfloor \frac{r}{j} \rfloor, \mathcal{R})\right)$$

this expected time accounts for averaging over the $r!$ permutations of the objects of \mathcal{R} . Now, the expected time to insert the r^{th} object of \mathcal{S} results from further averaging over the r -random samples of \mathcal{S} which yields :

$$O\left(\sum_{j=1}^r \frac{1}{r} f_0(\lfloor \frac{r}{j} \rfloor, \mathcal{S})\right).$$

The proof of Corollary 2.5 results from summing over r from 1 to n . □

Two immediate consequences of Theorem 2.4 are the following :

- If $f_0(x, \mathcal{S}) = \Theta(x)$, the space complexity is $O(n)$ and the time to insert a new object is $O(\log n)$.
- If $f_0(x, \mathcal{S}) = \Theta(x^\alpha)$ (with $\alpha \geq 1$), the space complexity is $O(n^\alpha)$ and the time to insert a new object is $O(n^{\alpha-1})$.

2.4 Removing the update conditions

In some cases the three update conditions can be removed. We will show in Section 3 that removing the update conditions (especially Condition 1) will lead, in some cases, to simpler algorithms.

Constant test time and linear update time

Update conditions 2 and 3 can be relaxed. If the time required to check if a region and an object are in conflict surpasses $O(1)$, the overcost will simply appear as a multiplicative factor in the overall complexity. If the time required by the creation substep surpasses a linear function of the number of conflicts, it is in general not difficult to charge the overcost to the overall complexity. This analysis has been done for example for the incremental construction of higher order Voronoï diagrams [BDT].

Bounded number of sons

It is more interesting to attempt to remove update condition 1. The preceding analysis works because each node in the I-DAG is associated to a region and all the relevant quantities can be expressed as functions of the number of nodes. If the condition is not fulfilled, we must count the number of edges of the I-DAG and to that purpose, we introduce the notion of *bicycles*. A bicycle is a pair of regions of $\mathcal{F}(\mathcal{S})$ occurring as a father and one of its sons in the I-DAG associated with at least one permutation of the object set \mathcal{S} .

Notice that the maximum number of objects defining a bicycle is at most $2b$ and thus is still bounded. An object is in conflict with a bicycle if it does not belong to the set of objects that define the bicycle and if it is in conflict with at least one of the two regions forming the bicycle. Thus the influence range of a bicycle is the union of the influence range of the two regions forming the bicycle, excepting the objects defining these regions.

In analogy with the notation used for regions, the additional notations $\mathcal{G}_j(\mathcal{S})$, $\mathcal{G}_{\leq j}^i(\mathcal{S}) \dots$ are naturally derived. We define $g_0(r, \mathcal{S})$ to be $E(|\mathcal{G}_0(\mathcal{R})|)$ the expected size of $\mathcal{G}_0(\mathcal{R})$ for r -random samples of \mathcal{S} .

With these definitions, the following lemma can be proved, using the random sampling technique, in a way similar to Lemma 2.1.

Lemma 2.6

$$|\mathcal{G}_{\leq j}^i(\mathcal{S})| = O\left(j^i g_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$$

We can now compute the expected storage required by the I-DAG :

Lemma 2.7 *If \mathcal{S} has cardinality n , the expected size of the I-DAG of \mathcal{S} is :*

$$O\left(\sum_{j=1}^n \frac{g_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})}{j}\right)$$

Proof. The size of the I-DAG is linearly related to the number of its edges. A necessary condition for a given bicycle G to occur as an edge in the I-DAG is that the i objects defining G are inserted before the j objects in conflict with G (i.e. in conflict with one of the two regions associated with G). So the probability that $G \in \mathcal{G}(\mathcal{S})$ arises in the I-DAG is less than $\frac{i!j!}{(i+j)!}$. Calculations analogous to those appearing in the proof of Lemma 2.2 yield the result. \square

Lemma 2.8 *The expected time for inserting the n^{th} object in the I-DAG is*

$$O\left(\frac{1}{n} \sum_{j=1}^n g_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$$

Proof. Let O be the n^{th} object to be inserted. The number of times a node is visited is equal to the number of its parents which are in conflict with O . Thus the number of performed tests is no more than the number of bicycles in conflict with O occurring in the I-DAG.

Let G be a bicycle of $\mathcal{G}_j^i(\mathcal{S})$. G is in conflict with O and occurs as an edge in the I-DAG if the following two necessary conditions are satisfied : 1) the i objects defining G are inserted before the j objects in conflict with G , 2) one of the j objects is O . The probability that G is in conflict with O is thus no more than $\frac{j}{n} \frac{i!(j-1)!}{(i+j-1)!}$. The result is then achieved as in Lemma 2.3. \square

Lemmas 2.7 and 2.8 prove the main result of this section :

Theorem 2.9 *If the set of already inserted objects \mathcal{S} has cardinality n and if update conditions 2 and 3 are fulfilled (but not update condition 1), the I-DAG of \mathcal{S} requires $O\left(\sum_{j=1}^n \frac{g_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})}{j}\right)$ expected memory space. The insertion of a new object can be done in $O\left(\frac{1}{n} \sum_{j=1}^n g_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$ expected update time.*

Corollary 2.10 *Under update conditions 2 and 3, the total expected time to build a I-DAG for a set \mathcal{S} of n object is : $O\left(\sum_{j=1}^n g_0(\lfloor \frac{n}{j} \rfloor, \mathcal{S})\right)$*

2.5 Faster object location

If the following additional property is verified, it is possible to get a better complexity result for the search of a single region in conflict with a new object.

Property 3 The influence range of the region associated to a node is included in the union of the ranges of its children.

If Property 3 holds, then a conflict with a given new object is found by a simple path from the root of the I-DAG to a leaf.

Theorem 2.11 *If Property 3 holds then a conflict with any new object can be found in $O(\log n)$ time provided that the n objects that were inserted in the I-DAG have been inserted in random order.*

Proof. Let o be the new object and F be a region on the path from the root to a leaf of the I-DAG in conflict with o . Suppose that F has zero width after the insertion of the k^{th} object. If F is defined by i objects, the conditional probability that F has been created during the insertion of the k^{th} object is $\frac{i}{k} \leq \frac{b}{k}$. Indeed, for F to be created at step k , the k^{th} inserted object must be one of the i objects defining F . (It is important to notice that the above probability is conditioned by the fact that the k first objects are given.)

Averaging this probability over the $\binom{n}{k}$ possible subsets of k objects introduced first in the I-DAG yields a probability less than $\frac{b}{k}$ that the node on the path would change after the insertion of the k^{th} object. Thus the number of visited nodes is less than $\sum \frac{b}{k} = O(\log n)$. \square

Let us make some remarks about Theorem 2.11. First it is important to notice that the cost studied here is expected over all possible orders to insert the objects in the I-DAG, but there is no hypothesis on the new object, by opposition to Lemma 2.3 where all objects, including the last one, are supposed to satisfy the randomization hypothesis. Secondly, in many applications, it is possible to find all conflicts with a new object from a single one in time proportional to the actual number of conflicts (for example by the use of some neighborhood notions). The faster location may be used as a first step of the insertion of a new object in the I-DAG. A last remark concerns the worst case ; though this article takes interest in randomized complexities, the faster location has a worst case running time $O(n)$ which is better than the general location step.

2.6 Queries

In some applications, queries consist of finding the regions having zero width which are in conflict with a given element of the object universe : this is just a special instance of a location substep. In such cases, the I-DAG can be used and the randomized analysis of Theorem 2.4 holds, provided that the query object

q together with the set of objects \mathcal{S} introduced in the I-DAG satisfy the randomization hypothesis, i.e. the $(n + 1)!$ permutations of $\{q\} \cup \mathcal{S}$ are likely to occur.

In other applications, regions and queries are such that the answer to a query consists of exactly one region of $\mathcal{F}_0(\mathcal{S})$ for any set \mathcal{S} . Such a query will be answered by a location substep that will traverse only one path from the root to a leaf. This yields the following strong variant of Lemma 2.3.

Theorem 2.12 *Assume that regions and queries are such that the answer to a query consists of exactly one region of $\mathcal{F}_0(\mathcal{S})$ for any set \mathcal{S} . Then any query can be answered in $O(\log n)$ time provided that the n objects that were inserted in the I-DAG have been inserted in random order.*

Proof. The proof is similar to the proof of Theorem 2.11. □

3 Applications

3.1 Convex hulls

We consider the geometric problem of computing the convex hull of a set of points. In the plane optimal deterministic on-line algorithms are known with time complexity $O(\log n)$ per update (see, for instance, [AES85]). In d -dimensional space, the worst case complexity of the convex hull of n points is $\Omega(n^{\lfloor \frac{d}{2} \rfloor})$. The best known deterministic algorithm, due to Seidel [Sei81, Ede87], requires $O(n \log n + n^{\lfloor \frac{d+1}{2} \rfloor})$ time and $O(n^{\lfloor \frac{d}{2} \rfloor})$ space. The algorithm is incremental but its complexity is amortized over the n insertions. This result is optimal in the worst-case for d larger than 2 and even. An output sensitive algorithm with time complexity $O(n^2 + f \log n)$ where f is the number of faces of the convex hull is reported in [Sei86].

Randomized algorithms have also been recently proposed by Clarkson and Shor [CS89] and by Mulmuley [Mul89]. These algorithms are incremental and optimal but static. A more on-line algorithm for $d \leq 3$ can be found in [GKS92]; however the analysis is only amortized over the n insertions.

We present in this section two (randomized) algorithms that are both on-line and optimal in any dimension. For brevity, we expose here only the two dimensional case. The extension to higher dimensional spaces is quite straightforward.

First algorithm

Objects are points of the plane. Regions are defined by three points. The region PQR associated with P , Q and R consists of the union of two half-planes : one

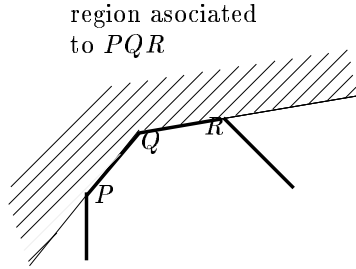


Figure 1: Definition of regions for the convex hull problem.

bounded by line (PQ) and not containing R and the other bounded by line (QR) and not containing P (see Figure 1). An object is in conflict with a region if and only if it lies in the region.

Now let P , Q and R be three points in \mathcal{S} , the set of points already inserted. The region associated with these points has zero width if and only if P , Q and R are three consecutive vertices of the convex hull. So computing the convex hull of \mathcal{S} is equivalent to computing the zero width regions.

Let us now describe the algorithm. Suppose that the I-DAG has been constructed for the points in \mathcal{S} and that we want to insert a new point M . The location substep gives the regions of $\mathcal{F}_0(\mathcal{S})$ containing M . If M belongs to the interior of the convex hull, there is no such region, $\mathcal{F}_0(\mathcal{S} \cup \{M\}) = \mathcal{F}_0(\mathcal{S})$ and the I-DAG is not modified. Otherwise, let P_1 and P_k be the two vertices adjacent to M in the new convex hull and P_2, \dots, P_{k-1} the chain of vertices which are no longer vertices of the convex hull after the insertion of M (see Figure 2). The regions of $\mathcal{F}_0(\mathcal{S})$ containing M are $P_{l-1}P_lP_{l+1}$, for $1 \leq l \leq k$. By a simple test on these regions, we can determine $P_0P_1P_2$ and $P_{k-1}P_kP_{k+1}$ (M belongs to only one of the two half-planes defining the region). The I-DAG is then modified in the following manner : the width of the selected regions is incremented and three new regions are created, namely

- P_0, P_1, M as a son of P_0, P_1, P_2
- M, P_k, P_{k+1} as a son of P_{k-1}, P_k, P_{k+1}
- P_1, M, P_k as a son of both P_0, P_1, P_2 and P_{k-1}, P_k, P_{k+1} .

It is clear that the properties of the I-DAG are preserved and that the update conditions are satisfied. Here $f_0(r, \mathcal{S})$ is the expected size of the convex hull of r points of \mathcal{S} which is clearly $O(r)$, so applying Theorem 2.4 we deduce :

Proposition 3.1 *The convex hull of n points in the plane can be computed on-line with $O(n)$ expected space and $O(\log n)$ expected update time.*

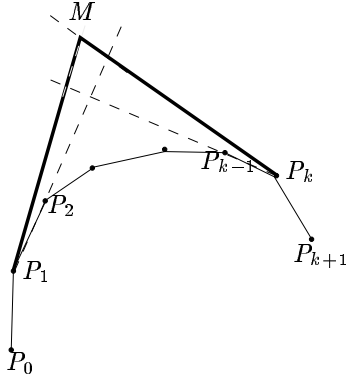


Figure 2: Inserting a new point in the convex hull

These results can be generalized to any dimension. The regions are defined by $d+1$ points and are unions of two half-spaces. The zero width regions correspond to $(d-2)$ -faces of the convex hull and their two adjacent $(d-1)$ -faces. The expected size $f_0(r, \mathcal{S})$ of the convex hull of r points is $O(r^{\lfloor \frac{d}{2} \rfloor})$.

Proposition 3.2 *The convex hull of n points in d -space can be computed on-line with $O(n^{\lfloor \frac{d}{2} \rfloor})$ expected space and $O(\log n)$ expected update time if $d \leq 3$ and $O(n^{\lfloor \frac{d}{2} \rfloor - 1})$ expected update time if $d > 3$.*

These results are optimal.

As far as queries are concerned, the above results show that one can decide if a point lies inside or outside the convex hull of n points in $O(\log n)$ expected time in the plane and $O(n^{\lfloor \frac{d}{2} \rfloor - 1})$ expected time in d -space.

Second algorithm

It might look more natural to take half-spaces as regions. In that case, as will become clear below, the number of sons is not bounded and update condition 1 is not satisfied. However the result of Section 2.4 proves that the resulting algorithm has the same complexity as the one above.

We describe the algorithm for the two dimensional case. It can be generalized to any dimension with no difficulty. \mathcal{O} still denotes the set of points of the plane. Regions are now defined by only two points. The regions PQ and QP are the two half planes limited by the line (PQ) . A point is in conflict with the region PQ if it lies inside the corresponding half-plane. If a region PQ has zero width, then $[PQ]$ is an edge of the convex hull.

In addition to the standard information stored in each node of the I-DAG, we also maintain at each leaf which is associated with an edge E of the convex hull two pointers towards the two leaves associated with the two edges of the convex hull adjacent to E . When a new point M is inserted, the location substep provides all the half-planes with current width 0 in conflict with M . These half-planes correspond to a chain of edges of the convex hull $P_1P_2, \dots, P_{k-1}P_k$ (see Figure 2). The two extremal edges P_1P_2 and $P_{k-1}P_k$ are identified, by testing if their two neighbours are not both in conflict with M . Two new regions P_1M and MP_k are created. In order to satisfy Property 2, P_0P_1 and P_1P_2 are made parents of P_1M ; similarly $P_{k-1}P_k$ and P_kP_{k+1} are made parents of MP_k . The neighborhood relationships are updated: P_0P_1 and P_1M become adjacent and similarly P_kP_{k+1} and MP_k , and P_1M and MP_k .

Notice that the width of P_0P_1 is still zero and that this region may have other sons in the future: update condition 1 is not satisfied.

As described in Section 2.4, we introduce the notion of a bicycle. Here a bicycle is defined by two regions PQ and QR sharing a point of definition (a bicycle here is a region of the previous section; see Figure 1). The zero width bicycles are of two kinds. The first ones are associated with two regions with zero width: P_0P_1 and P_1M in Figure 2. They correspond to two consecutive edges of the convex hull. The second ones are associated with a region of zero width and a region in conflict with the additional point of definition of the other: P_1P_2 and P_1M in Figure 2. They correspond to an edge E of the convex hull and to an edge E' , incident to one of the endpoints of E and whose supporting line separates a 1-set of \mathcal{S} . Using the results on the number of k -sets (see, for instance, [Ede87]) we conclude that $g_0(r)$ is $O(r^{\lfloor \frac{d}{2} \rfloor})$. Thus Theorem 2.9 implies that this simpler algorithm has the same complexity as the algorithm of the previous section.

3.2 Voronoi diagrams

Using the well known correspondence between Voronoi diagrams in d dimensions and convex hulls in $d + 1$ dimensions, we immediately deduce from the previous section two optimal on-line algorithms to construct the Voronoi diagrams of points in any dimension. A direct presentation that does not use this correspondence has already been described in detail in [BT].

Let us consider now the case of Voronoi diagrams of line segments in two dimensions (Figure 3). Here \mathcal{O} is the set of all line segments of the euclidean plane. Let P , Q , R and S be four segments and let Γ be the portion of the bisector of P and Q extending between the two points equidistant from P , Q , R and P , Q , S respectively (see Figure 4). The region associated with P , Q , R and S is the union of the disks tangent to P and Q whose centers lie on Γ . A line segment and a region are in conflict if and only if they intersect. A region has