Improved Incremental Randomized Delaunay Triangulation.*

Olivier Devillers†

1 Introduction

The computation of the Delaunay triangulation of a set of n points in the plane is one of the classical problems in computational geometry and plenty of algorithms have been proposed to solve it.

These Delaunay algorithms can have different characteristics:

- Optimal on worst case data, i.e. $O(n \log n)$ time.
- Good complexity on random data only
- Randomized
- On-line vs off-line

In the current trade-off between algorithmic simplicity, practical efficiency and theoretical optimality, practitioners often choose the simplicity and practical efficiency taking the risk of having bad performance on some special kind of data.

Our aim is to conciliate many of the above aspects, namely to obtain an incremental algorithm using simple data structure having good practical performance on realistic input and still provable $O(n \log n)$ computation time on any data set.

Previous related work

Our work is strongly related to some previous algorithms for Delaunay triangulation. All these algorithms are incremental and their complexity is randomized, they use some location structure to find where the new point is inserted, and then update the triangulation.

The first idea of a randomized incremental construction for the Delaunay triangulation [BT86] uses a location structure based on the history of the Delaunay triangulation: the Delaunay tree. Point p_i is inserted at time i, and to find where point p_n fell, p_n is located in all the triangulations at times 1 to $n-1$; the location at time $i+1$ is deduced from the location at time i. This idea yields an expected $O(n \log n)$ complexity [BT93, GKS92] if the points are inserted in a random order. The drawbacks of this approach are the following: the location structure consists of the history of the construction and thus strongly depends on the insertion order, and the additional memory needed cannot be controlled. (The expected memory is proved to be $O(n)$ and is experimentally about twice the size of the final triangulation.)

Mulmuley [Mul91] proposed a location structure independent of the insertion order. The structure has $O(\log n)$ levels, each level being a random sample of the level below. At each level, the Delaunay triangulation of the points is computed, and the overlapping triangles at different levels are linked to enable location of new points. This structure has the advantage of being independent of the order of insertion, of ensuring an $O(\log^2 n)$ location time for any point, and of allowing deletions in an easier way than the Delaunay tree [DMT92]. However, the additional memory is still important and the location structure is not especially simple.

In 1996, Mücke, Saias and Zhu [MSZ96] proposed a very simple structure to handle triangulation of random points. The structure reduces to a random subset of $\sqrt[3]{n}$ points, and pointers from these points to an incident triangle in the Delaunay triangulation. A new point is located by finding the nearest neighbor in the sample by brute force, and walking

*This work was partially supported by ESPRIT LTR 21957 (CGAL)
†INRIA, BP93, 06902 Sophia Antipolis. Olivier.Devillers@sophia.inria.fr.

SCG 98 Minneapolis Minnesota USA
Copyright ACM 1998 0-89791-973-4/98/ 6...$5.00

in the triangulation. For evenly distributed points, the expected complexity of the algorithm is $O(n^{\frac{4}{3}})$ with a small constant, which makes it competitive with many $O(n \log n)$ algorithms. But for some data (for example points on a parabola) the complexity increases to $O(n^{\frac{5}{3}})$.

Overview

Our approach uses a structure with levels similar to Mulmuley, but with simple relations between levels. This allows better control of the memory overhead. The transition between two levels is not direct as in Mulmuley, but uses a march similar Mücke, Saias and Zhu to locate point in triangulations.

In Section 2 we present the algorithm, in Section 3 we prove that the expected complexity of constructing the Delaunay triangulation is $O(n \log n)$. The parameters of the data structure are then tuned to minimize the constant in the case of random points and are shown to yield an excellent behavior in Section 4, we pay special attention to the comparison with the method of Mücke, Saias and Zhu. Finally we give some implementation remarks and practical results in Section 5.

2 Algorithm

Let $\mathcal{S}$ be a set of n sites in the plane. The aim is to compute the Delaunay triangulation $\mathcal{DT}_{\mathcal{S}}$ of $\mathcal{S}$ and to maintain it efficiently under insertions and deletions.

2.1 The location structure

The algorithm uses a data structure composed of different levels. Level i contains the Delaunay triangulation $\mathcal{DT}_i$ of a set of sites $\mathcal{S}_i$.

The sets $\mathcal{S}_i$ forms a decreasing sequence of random subsets of $\mathcal{S}$ based on a Bernoulli sampling technique [MR95, Mul94]:

$$\mathcal{S} = \mathcal{S}_0 \supseteq \mathcal{S}_1 \supseteq \mathcal{S}_2 \supseteq \ldots \supseteq \mathcal{S}_{k-1} \supseteq \mathcal{S}_k$$

$$Prob(p \in \mathcal{S}_{i+1} \mid p \in \mathcal{S}_i) = \frac{1}{\alpha} \in]0,1[.$$

The data structure is fairly simple: it contains the points of $\mathcal{S}$ and the triangles of all the triangulations $\mathcal{DT}_i$. A point $p \in \mathcal{S}$ such that $p \in \mathcal{S}_i \subseteq \ldots \subseteq \mathcal{S}_0$ and $p \notin \mathcal{S}_{i+1}$ is said to be a *vertex of level i* and has a link to a Delaunay triangle of $\mathcal{DT}_j$ incident to p for all j for $0 \leq j \leq i$. A triangle of $\mathcal{DT}_i$ has links to its three neighbors in $\mathcal{DT}_i$ and to its three vertices.The number k of levels is not fixed; for each point random trials decide its level, and the point with highest level determines k.

2.2 Location of a query

For the location of a query q, we start at a known vertex v_{k+1} of the highest level k. Then we search for v_k, the vertex of $\mathcal{DT}_k$ nearest to q. Since v_k is also a vertex of $\mathcal{DT}_{k-1}$, we search for v_{k-1}, the nearest neighbor of q in $\mathcal{DT}_{k-1}$, starting at v_k. The search is continued descending the different levels. At each level i, the nearest vertex v_i of q in $\mathcal{DT}_i$ is determined.

At level i the search of v_i is carried out in three phases:

- First phase: from v_{i+1}, we have a link to a triangle of $\mathcal{DT}_i$ having v_{i+1} as vertex. All triangles incident to v_{i+1} are explored to find the triangle containing the segment $v_{i+1}q$.
- Second phase: all the triangles of $\mathcal{DT}_i$ intersected by $v_{i+1}q$ are visited, walking along the segment $v_{i+1}q$ up to the triangle t_i that contains q.
- Third phase: using neighborhood relationships between triangles, we will traverse few triangles of $\mathcal{DT}_i$ from t_i to find v_i. If $vv'v''$ are the three vertices of t_i, and, without loss of generality, v is closer to q than v' and v'', then v_i is either v or it lies in the disk of center q and passing through v (shaded on Figure 1a); thus the search for v_i has to be done only in the direction of the neighbors of t_i through the edges vv' and vv'' and the neighbor through the edge $v'v''$ can be ignored (the portion of the shaded disk in that direction is inside the disk through $vv'v''$ which is empty) . For each such triangle, the distance to the new vertex is computed and the algorithm maintains the closest visited vertex. For a visited triangle $ww'w''$ such that w is the nearest to q among $ww'w''$ the neighbor triangle through edge ww' (resp ww'') will be visited if angle qww' is smaller than $\frac{\pi}{2}$ (Figure 1b).

Figure 1c show the triangles visited by the different phases of the search.

2.3 Updates

Because of its simplicity, the data structure is fairly easy to update. Maintaining it dynamically provides a fully dynamic triangulation algorithm. The links between the different levels do not use any complicated data structure simply vertices know a triangle at all levels in which they appear.

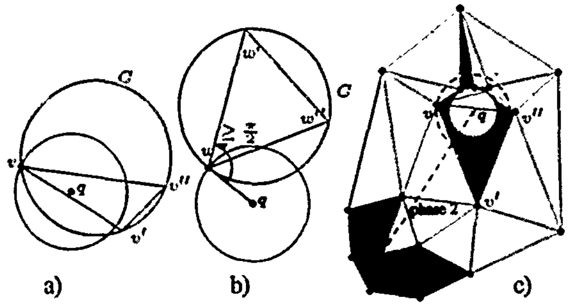

Figure 1: Search for v_i.

To delete a point from $\mathcal{S}$, just delete the corresponding vertex at all the levels where it appears, which can be done in time sensitive to d the degree of that vertex. On average $d = 6$ and thus some of the following algorithms can be used. A complicated algorithm [AGSS89] of deterministic complexity $O(d)$, a simple randomized $O(d)$ algorithm [Che86] can be used or simpler solutions of complexity $O(d \log d)$ or even $O(n^2)$ may be good in practice.

Inserting a point in $\mathcal{S}$ reduces to locating the new point at all levels, computing its level i and inserting the new vertex at all levels $j, 0 \leq j \leq i$ (which is sensitive to the degree of the new vertex once the location is done). The insertion using the standard algorithm [Law77].

3 Worst-case randomized analysis

The analysis will rely on the randomization in the construction of the random subsets $\mathcal{S}_i$ and the points of $\mathcal{S}$ are assumed to be inserted in a random order. In this section, no assumption applies to the data distribution, which can be in the worst case. As usual in theoretical computational geometry, we make only an asymptotic analysis and give rough upper bounds for the constants. In the next section, parameter α will be tuned to get a tight constant in the special case of evenly-distributed points.

Let $\mathcal{S}$ be a set of n points organized in the structure described in Section 2 and q a point to be inserted in $\mathcal{S}$. Since we have assumed a random insertion order, q is a random point of $\mathcal{S} \cup \{q\}$.

We denote $n_i = |\mathcal{S}_i|$ and $\mathcal{R}_i = \mathcal{S}_i \cup \{q\}$.

Notice that, thanks to the random insertion order, $\mathcal{R}_i$ is a random subset of size $n_i + 1$ of $\mathcal{R}_{i-1}$ and q is a random element of $\mathcal{R}_i$.

The cost of exploring all the triangles incident to v_{i+1} at the first phase of the march of level i is the degree of v_{i+1} in $\mathcal{DT}_i$. The cost of the second phase is the number of triangles intersected by segment $v_{i+1}q$. The cost of the third phase is the number of candidate vertices visited during the search of v_i from t_i.

Lemma 1 *The expected degree of v_i in $\mathcal{DT}_{i-1}$ is $O(1)$.*

Proof Let $\mathcal{NN}$ be the nearest neighbor graph of $\mathcal{R}_i$: that is, the vertices of $\mathcal{NN}$ are the points of $\mathcal{R}_i$, and $q, v \in \mathcal{R}_i$ define an edge of $\mathcal{NN}$ if and only if v is the nearest neighbor of q (denoted by $v = NN(q)$) or q is the nearest neighbor of v in $\mathcal{R}_i$. $\mathcal{NN}$ is well known to be a subgraph of $\mathcal{DT}_{\mathcal{R}_i}$, the Delaunay triangulation of $\mathcal{R}_i$, and to have maximum degree 6 [PY92].

We denote by $d^{\circ}_{\mathcal{DT}_{i-1}}(v)$ the degree of v in $\mathcal{DT}_{i-1}$, and by $E_{v \in \mathcal{R}_i \subset \{q\}}$ the expectation when v is chosen uniformly in $\mathcal{R}_i \subset \{q\}$. Then we have

$$E_{v \in \mathcal{R}_i \subset \{q\}} \left(d^{\circ}_{\mathcal{DT}_{i-1}}(v) \right)$$
$$= E_{v \in \mathcal{R}_{i-1} \subset \{q\}} \left(d^{\circ}_{\mathcal{DT}_{i-1}}(v) \right) < 6$$

notice that $d^{\circ}_{\mathcal{DT}_{i-1}}(v)$ is a random variable; result holds since $\mathcal{R}_i$ and $\mathcal{R}_{i-1} \subset \{q\}$ are random subsets of $\mathcal{R}_{i-1}$ and that the average degree of a vertex in a triangulation is less than 6.

But even if q is a random point in $\mathcal{R}_i$, the vertex v_i, the nearest neighbor of q in $\mathcal{R}_i$, is not uniformly random.

$$\begin{aligned}
&E_{q \in \mathcal{R}_i} \left(d^{\circ}_{\mathcal{DT}_{i-1}}(NN(q)) \right) \\
&= E \left(\frac{1}{|\mathcal{R}_i|} \sum_{q \in \mathcal{R}_i} d^{\circ}_{\mathcal{DT}_{i-1}}(NN(q)) \right) \\
&= \frac{1}{|\mathcal{R}_i|} E \left(\sum_{v \in \mathcal{R}_i} \sum_{q \in \{\rho; v = NN(\rho)\}} d^{\circ}_{\mathcal{DT}_{i-1}}(v) \right) \\
&< \frac{1}{|\mathcal{R}_i|} E \left(\sum_{v \in \mathcal{R}_i} 6 d^{\circ}_{\mathcal{DT}_{i-1}}(v) \right) \\
&\leq 36
\end{aligned}$$

■

Lemma 2 *Given $w \in \mathcal{R}_i$, the expected number of vertices q of $\mathcal{R}_i$ such that w belongs to the disk of center q and passing through the nearest neighbor of q in $\mathcal{R}_{i+1}$ is less than 6α.*

Proof Let $w \in \mathcal{R}_i$ and let $w = q_0, q_1, q_2 \dots q_k$ be the points of $\mathcal{R}_i$ lying in a section of angle $\frac{\pi}{3}$ having apex w sorted by increasing distance to w. Clearly, a disk of center q_l passing through q_j ($j < l$) cannot contain w and thus, if $q = q_l$, a necessary condition for w to be in the disk having as diameter the segment defined by q and the nearest neighbor of q in $\mathcal{R}_{i+1}$ is that no point of $\{q_0, \dots q_{l-1}\}$ is in the sample $\mathcal{R}_{i+1}$ which has probability $(1 - \frac{1}{\alpha})^l$.

Using six sections around w to cover the whole plane, and summing over the choice of $q \in \mathcal{R}_i$ we get the claimed result. Notice that the disk of center q and passing through the nearest neighbor of q contain the disk of diameter the line segment defined by these two points, and thus the bound apply also to that circle. ∎

Lemma 3 *The expected number of edges of $\mathcal{DT}_i$ intersecting segment qv_{i+1} is $O(\alpha)$.*

Proof Let e be an edge of $\mathcal{DT}_i$ intersecting segment qv_{i+1}. If e does not exist in $\mathcal{DT}_{\mathcal{R}_i}$, it means that e is an internal edge of the region retriangulated when q is inserted in $\mathcal{DT}_i$. Since q is a random point in $\mathcal{R}_i$, the expected number of such edges is 3 since it equals the average degree of q in $\mathcal{R}_i$ minus 3.

If e exists in $\mathcal{DT}_{\mathcal{R}_i}$, one end-point w of e must belong to the disk of diameter qv_{i+1}, denoted disk$[qv_{i+1}]$, (otherwise any disk through the end-points of e must contain q or v_{i+1} and e cannot belong to $\mathcal{DT}_{\mathcal{R}_i}$).

The expected number of edges of $\mathcal{DT}_{\mathcal{R}_i}$ intersecting disk$[qv_{i+1}]$ is bounded by the sum of the degrees of the vertices in disk$[qv_{i+1}]$

$$
\begin{aligned}
&E(\#\{e \in \mathcal{DT}_{\mathcal{R}_i} \text{ having an end-point} \in \mathcal{R}_i \cap \text{disk } [qv_{i+1}]\}) \\
&= \frac{1}{|\mathcal{R}_i|} \sum_{q \in \mathcal{R}_i} \sum_{w \in \mathcal{R}_i \cap \text{disk } [qv_{i+1}]} d^{\circ}_{\mathcal{DT}_{\mathcal{R}_i}}(w) \\
&= \frac{1}{|\mathcal{R}_i|} \sum_{w \in \mathcal{R}_i} d^{\circ}_{\mathcal{DT}_{\mathcal{R}_i}}(w) \, |\{q \in \mathcal{R}_i | w \in \text{disk } [qv_{i+1}]\}| \\
&\leq \frac{1}{|\mathcal{R}_i|} \sum_{w \in \mathcal{R}_i} d^{\circ}_{\mathcal{DT}_{\mathcal{R}_i}}(w) 6\alpha \quad \text{using Lemma 2} \\
&\leq 36\alpha \quad \text{using the bound of 6 on the average degree of } w
\end{aligned}
$$

Notice that Lemma 2 was established for a fixed w and a random q which allows to use it inside the sum over w. Thus we get a total expected cost for the march bounded by $36\alpha + 3$. ∎

Lemma 4 *The expected number of triangles of $\mathcal{DT}_i$ visited during the search for v_i from t_i is $O(\alpha)$.*

Proof All the triangles t examined in phase 3 have a vertex in the disk of center q passing through v_{i+1}. Thus we can argue similarly as in Lemma 3, denoting disk $|_c qv_{i+1}]$ the disk of center q through v_{i+1}:

$$
\begin{aligned}
&E(\#\{t \in \mathcal{DT}_{\mathcal{R}_i} \text{ having an end-point} \in \text{disk } |_c qv_{i+1}]\}) \\
&\leq \frac{1}{|\mathcal{R}_i|} \sum_{q \in \mathcal{R}_i} \sum_{w \in \mathcal{R}_i \cap \text{disk } |_c qv_{i+1}]} d^{\circ}_{\mathcal{DT}_{\mathcal{R}_i}}(w) \\
&\leq \frac{1}{|\mathcal{R}_i|} \sum_{w \in \mathcal{R}_i} d^{\circ}_{\mathcal{DT}_{\mathcal{R}_i}}(w) \, |\{q \in \mathcal{R}_i | w \in \mathcal{R}_i \cap \text{disk } |_c qv_{i+1}]\}| \\
&\leq \frac{1}{|\mathcal{R}_i|} \sum_{w \in \mathcal{R}_i} d^{\circ}_{\mathcal{DT}_{\mathcal{R}_i}}(w) 6\alpha \quad \text{using Lemma 2} \\
&\leq 36\alpha \quad \text{using the bound on the average degree of } w
\end{aligned}
$$

∎

Theorem 5 *The expected cost of inserting n^{th} point in the structure is $O(\alpha \log_\alpha n)$*

Proof By linearity of expectation, Lemmas 1, 3 and 4 prove that the expected cost at one level is $O(\alpha)$. Since the expected height of the structure is $\log_\alpha n$, we get the claimed result. (The analysis is similar to the ananlysis for skip lists [MR95].) ∎

Theorem 6 *The construction of the Delaunay triangulation of a set of n points is done in expected time $O(\alpha n \log_\alpha n)$ and $O(\frac{\alpha}{\alpha-1} n)$ space. The expectation is on the randomized sampling and the order of insertion, with no assumption on the point distribution.*

Proof Easy corollary of Theorem 5. ∎

4 Tuning parameters

We have proved that our structure is worst case optimal in the expected sense for any set of points. In this section, we will focus on more practical cases, and tune the algorithm to be optimal on random distribution. In that case, many events such as that a point has high degree and that it is the nearest neighbor of a random point can be considered as independent.

4.1 Phase 1

We can assume that, $d^{\circ}_{\mathcal{DT}_i}(v_{i+1}) = 6$ (and not only ≤ 36 as proved in Lemma 1). And thus if the turn around v_{i+1} is done in clockwise or counterclockwise direction depending on the position of segment $v_{i+1}q$ with respect to the starting

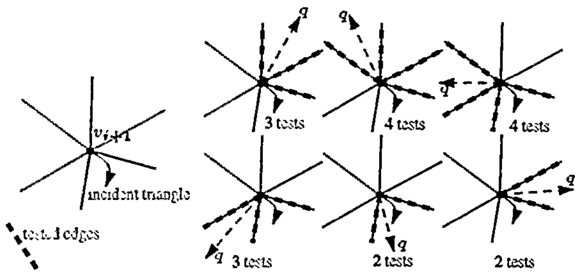

Figure 2: Number of orientation tests in phase 1

triangle, and assuming that this position is random around v_{i+1} the expected number of orientation tests is 3. Figure 2 shows the different cases to average, the edges $v_{i+1}w$ such that an orientation test $v_{i+1}wq$ is performed are indicated, for a typical degree 6 vertex in the triangulation.

4.2 Phase 2

Bose and Devroye [BD95] proved that the expected number of edges of a Delaunay triangulation of random points crossed by a line segment of length l is $O(l\sqrt{\gamma})$ where γ is the point density. Our experiments shows that the constant is 2.

The expected number of points in the disk of center q passing through v_{i+1} is $\alpha - 1$. Indeed, if the points of $\mathcal{R}_i$ are sorted by increasing distance from q, v_{i+1} is the first point in $\mathcal{R}_{i+1}$, thus the number of points in the disk is k with probability $(1-\frac{1}{\alpha})^k\frac{1}{\alpha}$, and the expected number is $\frac{1}{\alpha}\sum(1-\frac{1}{\alpha})^k = \alpha - 1$. Thus if l is the length of qv_{i+1} the density of points in $\mathcal{DT}_i$ is $\frac{\alpha}{\pi l^2}$.

Thus we conclude that the expected number of edges of $\mathcal{DT}_i$ intersecting segment qv_{i+1} is $2l\sqrt{\frac{\alpha}{\pi l^2}} = \frac{2\sqrt{\alpha}}{\sqrt{\pi}}$.

For each edge ww' crossed, two orientation tests are performed: if w is the newly examined vertex, orientations of triangles wqv_{i+1} and qww' are computed.

We have to point out, that in the orientation tests of kind wqv_{i+1}, the edge qv_{i+1} remains constant, and thus some computations do not need to be done for each test.

4.3 Phase 3

Phase 3 is more difficult to analyze precisely, but a rough bound is that the number of candidate vertices examined (with shortest distance) is less than two and that we examine less than 8 triangles in total.

In fact, we modified phase 3, instead of really searching for v_i, the nearest neighbor of q in $\mathcal{S}_i$, we just define v_i as the nearest among the three vertices of t_i. Thus this modified phase 3 reduced to three distance computations and two comparisons.

4.4 Tuning α

We will count more precisely the number of operation needed to evaluate our primitives. More exactly, we count the number of floating point operations (f.p.o.) without making diistinctions between additions, subtractions or multiplications.

The total evaluation at a given level is $3 + \frac{\sqrt{\alpha}}{\sqrt{\pi}}$ orientation tests involving qv_{i+1}, $\frac{\sqrt{\alpha}}{\sqrt{\pi}}$ other orientation tests and 3 distance computations.

Orientation tests always using points q and v_{i+1} can be done using 5 f.p.o. to initialize plus 4 f.p.o. for each test. Other orientation tests need 7 f.p.o. each, and square distance computations need 5 f.p.o. each.

Thus the total cost in terms of number of f.p.o. at level i is

$$5 + 4(3 + \frac{\sqrt{\alpha}}{\sqrt{\pi}}) + 7\frac{\sqrt{\alpha}}{\sqrt{\pi}} + 5 \cdot 3 = 32 + 6.2\sqrt{\alpha}.$$

Since the number of level is $\log_\alpha n = \frac{\log_2 n}{\log_2 \alpha}$ we get a cost of $c_0(n) = (32 + 6.2\sqrt{\alpha})\left\lceil\frac{\log_2 n}{\log_2 \alpha}\right\rceil$ which is close to its minimum ($\in [13.3\log_2 n, 14\log_2 n]$) for $\alpha \in [18, 90]$, with the minimum occuring for $\alpha \simeq 40$.

4.5 Comparison with [MSZ96]

Similar counting of f.p.o. in Mücke et al. algorithm, using a random sample of $\beta\sqrt[3]{n}$ points, produces a cost of

$$c_{MSZ}(n) = 5 + 4(3 + \frac{\sqrt{\frac{n}{\beta n^{\frac{1}{3}}}}}{\sqrt{\pi}}) + 7\frac{\sqrt{\frac{n}{\beta n^{\frac{1}{3}}}}}{\sqrt{\pi}} + 5\beta\sqrt[3]{n}$$

$$= 17 + \sqrt[3]{n}\left(\frac{6.2}{\sqrt{\beta}} + 5\beta\right)$$

which is close to its minimal value for $0.5 < \beta < 1$.

As shown by the comparison of the two curves in Figure 3, our method is potentially much better than [MSZ96], even for a small number of points. However, this method to analyze our approach hides the discontinuity of the cost, since the effective number of levels is necessarily an integer.

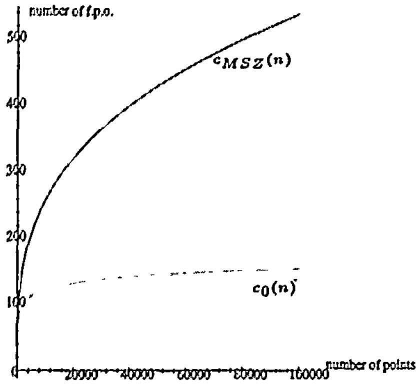

Figure 3: Comparison of number of floating point operations between $c_0(n)$ and $c_{MSZ}(n)$ for $\alpha = 40$ and $\beta = 1$.

To have a better comprehension of what happens for a small number of points, we can draw the cost of inserting a point in a structure having a fixed number of levels.

The classical walk from a random point in the structure costs

$$c_{walk}(n) = 5 + 4(3 + \frac{\sqrt{n}}{\sqrt{\pi}}) + 7\frac{\sqrt{n}}{\sqrt{\pi}} = 17 + 6.2\sqrt{n}$$

which is also the cost of inserting in our structure up to the time a second level is created.

When k levels have been created, the cost is

$$c_k(n) = c_{walk}\left(\frac{n}{\alpha^k}\right) + 15k + k \cdot c_{walk}(\alpha)$$

We can alternatively mix this multilevel approach with Mücke et al's, sampling at the first level of the structure. In that case, the cost is

$$c_k^*(n) = c_{MSZ}\left(\frac{n}{\alpha^k}\right) + 15k + k \cdot c_{walk}(\alpha)$$

This comparison (see Figure 4) shows that [MSZ96] ($c_1^*(n)$) becomes better than the simple march ($c_1(n)$) for $n > 40$. The two level structure ($c_2(n)$) becomes better than the single level structure ($c_1(n)$) for $n > 180$ and better than [MSZ96] ($c_1^*(n)$) for $n > 600$. The main information is that the structure presented in that paper should be significantly better than [MSZ96] for $10000 < n$.

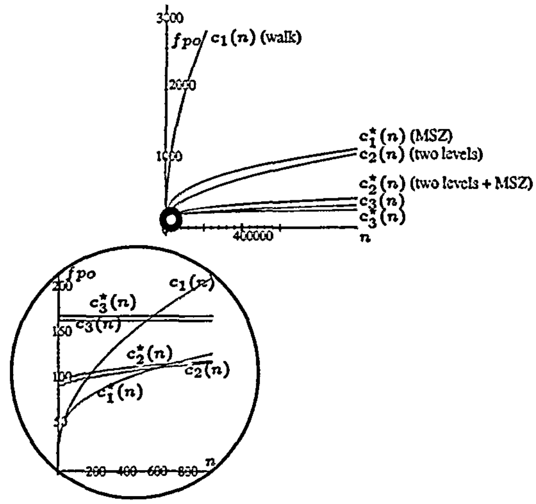

Figure 4: Comparison of number of floating point operations between $c_k(n)$ and $c_k^*(n)$ for $\alpha = 40$.

5 Implementation

5.1 Deletion

The above structure supports insertions and queries as explained above, but also deletions. Since there is no complicated data structure to maintain, deletions can be handled by just deleting the removed point at each level where it appears.

This can be done in output-sensitive time [Che87, AGSS89], and thus the deletion of a random point is done in expected constant time since a point appears at an expected constant number of levels and its expected degree k is also constant.

From a practical point of view, and to keep the simplicity of the algorithm, a simpler suboptimal algorithm should be preferred. It can be done in $O(k^2)$ time, for example by flipping to reduce the degree of the deleted vertex to 3, and flipping again to restore the Delaunay property. Another simple algorithm consists in finding the Delaunay triangle incident to an edge of the hole in $O(k)$ time which also yields an $O(k^2)$ time algorithm. Both algorithms are efficient in practice and needs only few micro-seconds (about 30 in a random triangulation) to delete a point once it had been localized.

5.2 Arithmetic degree

The algorithm above is designed to make a parsimonious use of high degree tests [TLP96]. More precisely, the location phase uses only orientation tests on three points in phases

1 and 2, and distance computation and angle comparisons with $\frac{\pi}{2}$ in phase 3. All these tests are degree 2 tests. Clearly, updates need to use in-circle tests which are of degree 4.

An alternative to phase 3 should have to use in-circle tests to limit the explored triangles in $\mathcal{DT}_i$ to those whose circumcircle contains q. Such variant may explore fewer triangles and be easier to analyze, but may use more degree 4 tests.

5.3 Robustness issues and degeneracies

Degeneracies are solved by handling special cases: if two points have the same coordinates, then the insertion is not done, if four points are cocircular, then the last point inserted is considered as inside the disk defined by the others.

We use exact arithmetic for 24 bits integers, and thus coordinates of our points are integers in range $[-16777216, 16777216]$ (up to a multiplication by a power of 2). Using this restricted kind of data, double precision computation is exact on degree 2 tests and almost never leads to precision problems on degree 4 predicates. Nevertheless, the exactness of all computations are verified by an arithmetic filter and exact computation is performed if needed.

5.4 Code parameters

The following parameters can be specified:

- maximal number of levels
- α the ratio between two levels
- the minimal number of points to use the higher level for point location
- the minimal number of points to use MSZ sampling at one of the higher levels
- β the constant for the size of MSZ sample.

Our default parameters are

- number of levels unlimited
- $\alpha = 30$.
- minimal size to use hierarchy is 20.
- minimal size to use MSZ is 20.
- $\beta = 1$.

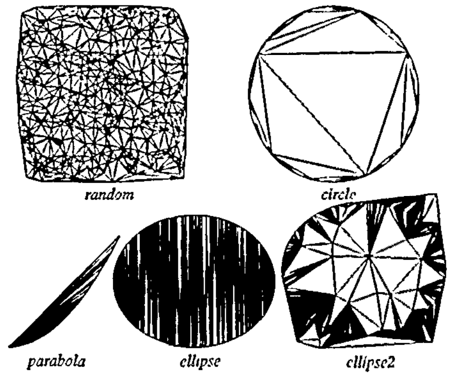

Figure 5: Data sets.

We found that the code is relatively insensitive to the parameters. For reasonable changes of these parameters, (up to a factor 2) the computation time is not greatly affected. Using these configuration parameters, our code can be used to run

- the usual walk algorithm (only one level and minimal size for MSZ=∞),
- the Mücke et al. algorithm [MSZ96] (only one level),
- the hierarchical algorithm described in this paper (minimal size for MSZ=∞),
- the mixed method suggested in Section 4.5 (default parameters above).

5.5 Experimental results

5.5.1 Data sets

We claim that our algorithm performs well on random point sets, and has acceptable worse case complexity. To illustrate this fact, we will test it with the realistic and degenerate data sets. For each kind of data, we used sets of size 5,000, 50,000 and 500,000 points. The coordinates are random on 24 bits and the constraints such that the points are on a parabola are verified, up to the rounding arithmetic errors.

- *random*: points evenly distributed in a square.
- *ellipse*: points evenly distributed on an ellipse.
- *ellipse2*: 95% points evenly distributed on an ellipse plus 5% points evenly distributed in a square.
- *circle*: points evenly distributed on a circle.
- *parabola*: points evenly distributed on a parabola,

distribution	size	walk	[MSZ96]	hierarchy	hierarchy + MSZ
random	5000	0.3	0.17	0.15	0.14
random	50000	12	3.8	2.7	2.3
random	500000	460	72	36	31
ellipse2	5000	0.53	0.34	0.21	0.20
ellipse2	50000	49	21	3.9	3.5
ellipse2	500000	930	760	57	49
ellipse	5000	2.2	0.46	0.31	0.21
ellipse	50000	187	21	3.9	3.7
ellipse	500000	long	270	54	55
parabola	5000	2.5	0.31	0.21	0.16
parabola	50000	87	5.9	3.2	3.0
parabola	500000	long	74	69	45
circle	5000	0.15	0.13	0.13	0.14
circle	50000	2.4	2.6	2.4	2.4
circle	500000	39	44	36	36

Figure 6: Running times

If the *circle* and *parabola* examples can be considered as pathological inputs, the *ellipse* and *ellipse2* examples are more realistic, Delaunay triangulation of points distributed on a curve occurs in practical applications, for example in shape reconstruction (see Figure 5).

5.5.2 Results

Following results are obtained on a Sun-Ultra1 200 MHz. The code is written in C++ and compiled with AT-T compiler with optimizing options. Time has been obtained with the `clock` command and is given in seconds. The time which is measured is just the Delaunay computation; it does not take into account the time for input or output.

Figure 6 gives the computation times for execution of the code with the different parameters described in Section 5.4. Since it is the same code, the low level primitives such as incircle tests or the walk in the triangulation are identical and it provides a fair comparison between the different methods.

The last column is always the fastest method. It is significantly better than MSZ for very large sets of random points, and the difference is even more important on data set *ellipse2* which is representative of real applications.

5.5.3 Comparison with other software

We have compared with some Delaunay softwares available on the WWW:

- `qhull` by Bradford Barber and Hannu Huhdanpaa, duality with 3D convex hull [BDH93] (available at http://www.geom.umn.edu/locate/qhull).
- `div-conquer` by Jonathan Shewchuk, divide and conquer [She96]
- `sweep` by Jonathan Shewchuk, plane sweep
- `incremental` by Jonathan Shewchuk, incremental with Mücke et al. localization.
  These three codes supports exact arithmetic on `double` (available at http://www.cs.cmu.edu/~quake/triangle.research.html).
- `Dtree` Delaunay tree structure[BT93] (time includes input) (available at http://www.inria.fr/prisme/logiciel/del-tree.html).
- `hierarchy` this paper, mixed with MSZ.

The execution times in seconds are in Figure 7. Our method is significantly faster than the other incremental method, especially in the ellipse cases. Our method is about 50% slower than the divide and conquer algorithm.

6 Conclusion

We proposed a new hierarchical data structure to compute the Delaunay triangulation of a set of points in the plane. It combines good worst case randomized complexity, fast behavior on real data, small memory occupation and dynamic updates (insertion and deletion of points).

Referring to Su and Drysdale [SD97] study of several techniques and our comparisons with Shewchuk implementation [She96] of some of these techniques, we have

distribution	size	qhull	sweep	div-conq	incr	Dtree	hier.
random	5000	0.65	0.21	0.11	0.29	1.4	0.14
random	50000	8.0	3.6	1.6	6.6	17	2.3
random	500000	101	53	22	150	swap	31
ellipse2	5000	0.54	0.21	0.13	0.75	1.3	0.20
ellipse2	50000	7.8	3.2	2.16	42	16	3.5
ellipse2	500000	420	46	29	2100	swap	49
ellipse	5000	0.83	0.18	0.14	2.1	1.3	0.21
ellipse	50000	57	2.8	2.4	110	14	3.7
ellipse	500000	swap	39	33	1400	swap	55
parabola	5000	3.9	0.16	0.11	2.0	1.2	0.16
parabola	50000	790	2.7	2.0	110	14	3.0
parabola	500000	swap	39	28	1800	swap	45
circle	5000	93	0.17	0.17	0.52	1.4	0.14
circle	50000	220	3.1	1.8	11	15	2.4
circle	500000	swap	22	43	240	swap	36

Figure 7: Comparisons with other softwares

shown that our implementation is competitive with other approaches on random data. Furthermore, we can prove that the performances remains good on pathological inputs. Finally, one of the main advantage of this algorithm is to allow a dynamic setting.

The main idea of our structure is to perform point location using several levels. The lowest level just consists of the triangulation, then each level contains the triangulation of a small sample of the levels below. Point location is done by marching in a triangulation to determine the nearest neighbor of the query at that level, then the march restart from that neighbor at the level below. Location at highest level is done using [MSZ96] which is efficient for small set of points.

One characteristics of the structure is that best time performance is obtained with a ratio of about three per cent between two levels, which yields to few levels (three or four typically) and a small memory occupation. The structure is simple and does not need additional features such as buckets.

Such structure can be generalized to other problems. The two main ingredients of the proofs are bounds on the maximal degree of the nearest neighbor graph and the expected degree of a random vertex in the Delaunay triangulation. The first generalizes well in higher dimension, while the second becomes an data sensitive parameter (constant for random points, $n^{\lceil (d-1)/2 \rceil}$ in the worst case). A generalization for computing the trapezoidal map can also be done.

Code

A demo version compiled for Sun Solaris and SGI is available at

http://www.inria.fr/prisme/logiciels/del-hierarchy/.

Acknowledgement

The author would like to thank Hervé Brönnimann, Jonathan Shewchuk, Jack Snoeyink and Mariette Yvinec for helpful discussions and careful reading of this paper.

References

[AGSS89] A. Aggarwal, L. J. Guibas, J. Saxe, and P. W. Shor. A linear-time algorithm for computing the Voronoi diagram of a convex polygon. *Discrete Comput. Geom.*, 4(6):591–604, 1989.

[BD95] P. Bose and L. Devroye. Intersections with random geometric objects. Technical report, School of Computer Science, McGill University, 1995. Manuscript.

[BDH93] C. B. Barber, D. P. Dobkin, and H. Huhdanpaa. The Quickhull algorithm for convex hull. Technical Report GCG53, Geometry Center, Univ. of Minnesota, July 1993.

[BT86] J.-D. Boissonnat and M. Teillaud. A hierarchical representation of objects: The Delaunay tree. In *Proc. 2nd Annu. ACM Sympos. Comput. Geom.*, pages 260–268, 1986.

[BT93] J.-D. Boissonnat and M. Teillaud. On the randomized construction of the Delaunay tree. *Theoret. Comput. Sci.*, 112:339–354, 1993.

[Che86] L. P. Chew. Building Voronoi diagrams for convex polygons in linear expected time. Technical Report PCS-TR90-147, Dept. Math. Comput. Sci., Dartmouth College, Hanover, NH, 1986.

[Che87] L. P. Chew. Constrained Delaunay triangulations. In *Proc. 3rd Annu. ACM Sympos. Comput. Geom.*, pages 215–222, 1987.

[DMT92] O. Devillers, S. Meiser, and M. Teillaud. Fully dynamic Delaunay triangulation in logarithmic expected time per operation. *Comput. Geom. Theory Appl.*, 2(2):55–80, 1992.

[GKS92] L. J. Guibas, D. E. Knuth, and M. Sharir. Randomized incremental construction of Delaunay and Voronoi diagrams. *Algorithmica*, 7:381–413, 1992.

[Law77] C. L. Lawson. Software for C^1 surface interpolation. In J. R. Rice, editor, *Math. Software III*, pages 161–194. Academic Press, New York, NY, 1977.

[MR95] R. Motwani and P. Raghavan. *Randomized Algorithms*. Cambridge University Press, New York, NY, 1995.

[MSZ96] Ernst P. Mücke, Isaac Saias, and Binhai Zhu. Fast randomized point location without preprocessing in two- and three-dimensional Delaunay triangulations. In *Proc. 12th Annu. ACM Sympos. Comput. Geom.*, pages 274–283, 1996.

[Mul91] K. Mulmuley. Randomized multidimensional search trees: Dynamic sampling. In *Proc. 7th Annu. ACM Sympos. Comput. Geom.*, pages 121–131, 1991.

[Mul94] K. Mulmuley. *Computational Geometry: An Introduction Through Randomized Algorithms*. Prentice Hall, Englewood Cliffs, NJ, 1994.

[PY92] M. S. Paterson and F. F. Yao. On nearest-neighbor graphs. In *Proc. 19th Internat. Colloq. Automata Lang. Program.*, volume 623 of *Lecture Notes Comput. Sci.*, pages 416–426. Springer-Verlag, 1992.

[SD97] P. Su and R. Drysdale. A comparison of sequential Delaunay triangulation algorithms. *Comput. Geom. Theory Appl.*, 7:361–386, 1997.

[She96] J. R. Shewchuk. Triangle: engineering a 2d quality mesh generator and Delaunay triangulator. In *First Workshop on Applied Computational Geometry*. Association for Computing Machinery, May 1996.

[TLP96] R. Tamassia, G. Liotta, and F. P. Preparata. Robust proximity queries in implicit Voronoi diagrams. In *Proc. 8th Canad. Conf. Comput. Geom.*, page 1, 1996.